# Linux Filesystems Documentation

***Release 6.8.0***

**The kernel development community**

**Jan 16, 2026**

# CONTENTS

This under-development manual will, some glorious day, provide comprehensive information on how the Linux virtual filesystem (VFS) layer works, along with the filesystems that sit below it. For now, what we have can be found below.

CHAPTER

ONE

# CORE VFS DOCUMENTATION

See these manuals for documentation about the VFS layer itself and how its algorithms work.

## 1.1 Overview of the Linux Virtual File System

Original author: Richard Gooch <rgooch@atnf.csiro.au>



### 1.1.1 Introduction

The Virtual File System (also known as the Virtual Filesystem Switch) is the software layer in the kernel that provides the filesystem interface to userspace programs. It also provides an abstraction within the kernel which allows different filesystem implementations to coexist.

VFS system calls open(2), stat(2), read(2), write(2), chmod(2) and so on are called from a process context. Filesystem locking is described in the document *Locking*.

#### Directory Entry Cache (dcache)

The VFS implements the open(2), stat(2), chmod(2), and similar system calls. The pathname argument that is passed to them is used by the VFS to search through the directory entry cache (also known as the dentry cache or dcache). This provides a very fast look-up mechanism to translate a pathname (filename) into a specific dentry. Dentries live in RAM and are never saved to disc: they exist only for performance.

The dentry cache is meant to be a view into your entire filespace. As most computers cannot fit all dentries in the RAM at the same time, some bits of the cache are missing. In order to resolve your pathname into a dentry, the VFS may have to resort to creating dentries along the way, and then loading the inode. This is done by looking up the inode.

#### The Inode Object

An individual dentry usually has a pointer to an inode. Inodes are filesystem objects such as regular files, directories, FIFOs and other beasts. They live either on the disc (for block device filesystems) or in the memory (for pseudo filesystems). Inodes that live on the disc are copied into the memory when required and changes to the inode are written back to disc. A single inode can be pointed to by multiple dentries (hard links, for example, do this).

To look up an inode requires that the VFS calls the lookup() method of the parent directory inode. This method is installed by the specific filesystem implementation that the inode lives in. Once the VFS has the required dentry (and hence the inode), we can do all those boring things like open(2) the file, or stat(2) it to peek at the inode data. The stat(2) operation is fairly simple: once the VFS has the dentry, it peeks at the inode data and passes some of it back to userspace.

#### The File Object

Opening a file requires another operation: allocation of a file structure (this is the kernel-side implementation of file descriptors). The freshly allocated file structure is initialized with a pointer to the dentry and a set of file operation member functions. These are taken from the inode data. The open() file method is then called so the specific filesystem implementation can do its work. You can see that this is another switch performed by the VFS. The file structure is placed into the file descriptor table for the process.

Reading, writing and closing files (and other assorted VFS operations) is done by using the userspace file descriptor to grab the appropriate file structure, and then calling the required file structure method to do whatever is required. For as long as the file is open, it keeps the dentry in use, which in turn means that the VFS inode is still in use.

### 1.1.2 Registering and Mounting a Filesystem

To register and unregister a filesystem, use the following API functions:

```
#include <linux/fs.h>

extern int register_filesystem(struct file_system_type *);
extern int unregister_filesystem(struct file_system_type *);
```

The passed struct file_system_type describes your filesystem. When a request is made to mount a filesystem onto a directory in your namespace, the VFS will call the appropriate mount() method for the specific filesystem. New vfsmount referring to the tree returned by ->mount() will be attached to the mountpoint, so that when pathname resolution reaches the mountpoint it will jump into the root of that vfsmount.

You can see all filesystems that are registered to the kernel in the file /proc/filesystems.

### struct file_system_type

This describes the filesystem. The following members are defined:

```
struct file_system_type {
        const char *name;
        int fs_flags;
        int (*init_fs_context)(struct fs_context *);
        const struct fs_parameter_spec *parameters;
        struct dentry *(*mount) (struct file_system_type *, int,
                const char *, void *);
        void (*kill_sb) (struct super_block *);
        struct module *owner;
        struct file_system_type * next;
        struct hlist_head fs_supers;

        struct lock_class_key s_lock_key;
        struct lock_class_key s_umount_key;
        struct lock_class_key s_vfs_rename_key;
        struct lock_class_key s_writers_key[SB_FREEZE_LEVELS];

        struct lock_class_key i_lock_key;
        struct lock_class_key i_mutex_key;
        struct lock_class_key invalidate_lock_key;
        struct lock_class_key i_mutex_dir_key;
};
```

**name**
the name of the filesystem type, such as ”ext2”, ”iso9660”, ”msdos” and so on

**fs_flags**
various flags (i.e. FS_REQUIRES_DEV, FS_NO_DCACHE, etc.)

**init_fs_context**
Initializes 'struct fs_context' ->ops and ->fs_private fields with filesystem-specific data.

**parameters**
Pointer to the array of filesystem parameters descriptors 'struct fs_parameter_spec'. More info in *Filesystem Mount API*.

**mount**
the method to call when a new instance of this filesystem should be mounted

**kill_sb**
the method to call when an instance of this filesystem should be shut down

**owner**
for internal VFS use: you should initialize this to THIS_MODULE in most cases.

**next**
for internal VFS use: you should initialize this to NULL

**fs_supers**

for internal VFS use: hlist of filesystem instances (superblocks)

s_lock_key, s_umount_key, s_vfs_rename_key, s_writers_key, i_lock_key, i_mutex_key, invalidate_lock_key, i_mutex_dir_key: lockdep-specific

The mount() method has the following arguments:

**struct file_system_type *fs_type**
describes the filesystem, partly initialized by the specific filesystem code

**int flags**
mount flags

**const char *dev_name**
the device name we are mounting.

**void *data**
arbitrary mount options, usually comes as an ASCII string (see "Mount Options" section)

The mount() method must return the root dentry of the tree requested by caller. An active reference to its superblock must be grabbed and the superblock must be locked. On failure it should return ERR_PTR(error).

The arguments match those of mount(2) and their interpretation depends on filesystem type. E.g. for block filesystems, dev_name is interpreted as block device name, that device is opened and if it contains a suitable filesystem image the method creates and initializes struct super_block accordingly, returning its root dentry to caller.

->mount() may choose to return a subtree of existing filesystem - it doesn't have to create a new one. The main result from the caller's point of view is a reference to dentry at the root of (sub)tree to be attached; creation of new superblock is a common side effect.

The most interesting member of the superblock structure that the mount() method fills in is the "s_op" field. This is a pointer to a "struct super_operations" which describes the next level of the filesystem implementation.

Usually, a filesystem uses one of the generic mount() implementations and provides a fill_super() callback instead. The generic variants are:

**mount_bdev**
mount a filesystem residing on a block device

**mount_nodev**
mount a filesystem that is not backed by a device

**mount_single**
mount a filesystem which shares the instance between all mounts

A fill_super() callback implementation has the following arguments:

**struct super_block *sb**
the superblock structure. The callback must initialize this properly.

**void *data**
arbitrary mount options, usually comes as an ASCII string (see "Mount Options" section)

**int silent**
whether or not to be silent on error

### 1.1.3 The Superblock Object

A superblock object represents a mounted filesystem.

#### struct super_operations

This describes how the VFS can manipulate the superblock of your filesystem. The following members are defined:

```
struct super_operations {
        struct inode *(*alloc_inode)(struct super_block *sb);
        void (*destroy_inode)(struct inode *);
        void (*free_inode)(struct inode *);

        void (*dirty_inode) (struct inode *, int flags);
        int (*write_inode) (struct inode *, struct writeback_control *wbc);
        int (*drop_inode) (struct inode *);
        void (*evict_inode) (struct inode *);
        void (*put_super) (struct super_block *);
        int (*sync_fs)(struct super_block *sb, int wait);
        int (*freeze_super) (struct super_block *sb,
                             enum freeze_holder who);
        int (*freeze_fs) (struct super_block *);
        int (*thaw_super) (struct super_block *sb,
                             enum freeze_wholder who);
        int (*unfreeze_fs) (struct super_block *);
        int (*statfs) (struct dentry *, struct kstatfs *);
        int (*remount_fs) (struct super_block *, int *, char *);
        void (*umount_begin) (struct super_block *);

        int (*show_options)(struct seq_file *, struct dentry *);
        int (*show_devname)(struct seq_file *, struct dentry *);
        int (*show_path)(struct seq_file *, struct dentry *);
        int (*show_stats)(struct seq_file *, struct dentry *);

        ssize_t (*quota_read)(struct super_block *, int, char *, size_t, loff_
↪t);
        ssize_t (*quota_write)(struct super_block *, int, const char *, size_t,
↪ loff_t);
        struct dquot **(*get_dquots)(struct inode *);

        long (*nr_cached_objects)(struct super_block *,
                             struct shrink_control *);
        long (*free_cached_objects)(struct super_block *,
                             struct shrink_control *);
};
```

All methods are called without any locks being held, unless otherwise noted. This means that most methods can block safely. All methods are only called from a process context (i.e. not from an interrupt handler or bottom half).

**alloc_inode**
this method is called by alloc_inode() to allocate memory for struct inode and initialize it. If this function is not defined, a simple 'struct inode' is allocated. Normally alloc_inode will be used to allocate a larger structure which contains a 'struct inode' embedded within it.

**destroy_inode**
this method is called by destroy_inode() to release resources allocated for struct inode. It is only required if ->alloc_inode was defined and simply undoes anything done by ->alloc_inode.

**free_inode**
this method is called from RCU callback. If you use call_rcu() in ->destroy_inode to free 'struct inode' memory, then it's better to release memory in this method.

**dirty_inode**
this method is called by the VFS when an inode is marked dirty. This is specifically for the inode itself being marked dirty, not its data. If the update needs to be persisted by fdatasync(), then I_DIRTY_DATASYNC will be set in the flags argument. I_DIRTY_TIME will be set in the flags in case lazytime is enabled and struct inode has times updated since the last ->dirty_inode call.

**write_inode**
this method is called when the VFS needs to write an inode to disc. The second parameter indicates whether the write should be synchronous or not, not all filesystems check this flag.

**drop_inode**
called when the last access to the inode is dropped, with the inode->i_lock spinlock held.

This method should be either NULL (normal UNIX filesystem semantics) or "generic_delete_inode" (for filesystems that do not want to cache inodes - causing "delete_inode" to always be called regardless of the value of i_nlink)

The "generic_delete_inode()" behavior is equivalent to the old practice of using "force_delete" in the put_inode() case, but does not have the races that the "force_delete()" approach had.

**evict_inode**
called when the VFS wants to evict an inode. Caller does *not* evict the pagecache or inode-associated metadata buffers; the method has to use truncate_inode_pages_final() to get rid of those. Caller makes sure async writeback cannot be running for the inode while (or after) ->evict_inode() is called. Optional.

**put_super**
called when the VFS wishes to free the superblock (i.e. unmount). This is called with the superblock lock held

**sync_fs**
called when VFS is writing out all dirty data associated with a superblock. The second parameter indicates whether the method should wait until the write out has been completed. Optional.

**freeze_super**
Called instead of ->freeze_fs callback if provided. Main difference is that ->freeze_super is called without taking down_write(&sb->s_umount). If filesystem implements it and wants ->freeze_fs to be called too, then it has to call ->freeze_fs explicitly from this callback. Optional.

**freeze_fs**
called when VFS is locking a filesystem and forcing it into a consistent state. This method is currently used by the Logical Volume Manager (LVM) and ioctl(FIFREEZE). Optional.

**thaw_super**
called when VFS is unlocking a filesystem and making it writable again after ->freeze_super. Optional.

**unfreeze_fs**
called when VFS is unlocking a filesystem and making it writable again after ->freeze_fs. Optional.

**statfs**
called when the VFS needs to get filesystem statistics.

**remount_fs**
called when the filesystem is remounted. This is called with the kernel lock held

**umount_begin**
called when the VFS is unmounting a filesystem.

**show_options**
called by the VFS to show mount options for /proc/<pid>/mounts and /proc/<pid>/mountinfo. (see ”Mount Options” section)

**show_devname**
Optional. Called by the VFS to show device name for /proc/<pid>/{mounts,mountinfo,mountstats}. If not provided then '(struct mount).mnt_devname' will be used.

**show_path**
Optional. Called by the VFS (for /proc/<pid>/mountinfo) to show the mount root dentry path relative to the filesystem root.

**show_stats**
Optional. Called by the VFS (for /proc/<pid>/mountstats) to show filesystem-specific mount statistics.

**quota_read**
called by the VFS to read from filesystem quota file.

**quota_write**
called by the VFS to write to filesystem quota file.

**get_dquots**
called by quota to get 'struct dquot' array for a particular inode. Optional.

**nr_cached_objects**
called by the sb cache shrinking function for the filesystem to return the number of freeable cached objects it contains. Optional.

**free_cache_objects**
called by the sb cache shrinking function for the filesystem to scan the number of objects indicated to try to free them. Optional, but any filesystem implementing this method needs to also implement ->nr_cached_objects for it to be called correctly.

We can't do anything with any errors that the filesystem might encountered, hence the void return type. This will never be called if the VM is trying to reclaim under GFP_NOFS conditions, hence this method does not need to handle that situation itself.

Implementations must include conditional reschedule calls inside any scanning loop that is done. This allows the VFS to determine appropriate scan batch sizes without having to worry about whether implementations will cause holdoff problems due to large scan batch sizes.

Whoever sets up the inode is responsible for filling in the "i_op" field. This is a pointer to a "struct inode_operations" which describes the methods that can be performed on individual inodes.

#### struct xattr_handler

On filesystems that support extended attributes (xattrs), the s_xattr superblock field points to a NULL-terminated array of xattr handlers. Extended attributes are name:value pairs.

**name**
Indicates that the handler matches attributes with the specified name (such as "system.posix_acl_access"); the prefix field must be NULL.

**prefix**
Indicates that the handler matches all attributes with the specified name prefix (such as "user."); the name field must be NULL.

**list**
Determine if attributes matching this xattr handler should be listed for a particular dentry. Used by some listxattr implementations like generic_listxattr.

**get**
Called by the VFS to get the value of a particular extended attribute. This method is called by the getxattr(2) system call.

**set**
Called by the VFS to set the value of a particular extended attribute. When the new value is NULL, called to remove a particular extended attribute. This method is called by the setxattr(2) and removexattr(2) system calls.

When none of the xattr handlers of a filesystem match the specified attribute name or when a filesystem doesn't support extended attributes, the various `*xattr(2)` system calls return -EOPNOTSUPP.

### 1.1.4 The Inode Object

An inode object represents an object within the filesystem.

#### struct inode_operations

This describes how the VFS can manipulate an inode in your filesystem. As of kernel 2.6.22, the following members are defined:

```
struct inode_operations {
        int (*create) (struct mnt_idmap *, struct inode *,struct dentry *,␣
↪umode_t, bool);
        struct dentry * (*lookup) (struct inode *,struct dentry *, unsigned␣
↪int);
```

```
        int (*link) (struct dentry *,struct inode *,struct dentry *);
        int (*unlink) (struct inode *,struct dentry *);
        int (*symlink) (struct mnt_idmap *, struct inode *,struct dentry *,
↪const char *);
        int (*mkdir) (struct mnt_idmap *, struct inode *,struct dentry *,umode_
↪t);
        int (*rmdir) (struct inode *,struct dentry *);
        int (*mknod) (struct mnt_idmap *, struct inode *,struct dentry *,umode_
↪t,dev_t);
        int (*rename) (struct mnt_idmap *, struct inode *, struct dentry *,
                       struct inode *, struct dentry *, unsigned int);
        int (*readlink) (struct dentry *, char __user *,int);
        const char *(*get_link) (struct dentry *, struct inode *,
                                 struct delayed_call *);
        int (*permission) (struct mnt_idmap *, struct inode *, int);
        struct posix_acl * (*get_inode_acl)(struct inode *, int, bool);
        int (*setattr) (struct mnt_idmap *, struct dentry *, struct iattr *);
        int (*getattr) (struct mnt_idmap *, const struct path *, struct kstat␣
↪*, u32, unsigned int);
        ssize_t (*listxattr) (struct dentry *, char *, size_t);
        void (*update_time)(struct inode *, struct timespec *, int);
        int (*atomic_open)(struct inode *, struct dentry *, struct file *,
                           unsigned open_flag, umode_t create_mode);
        int (*tmpfile) (struct mnt_idmap *, struct inode *, struct file *,␣
↪umode_t);
        struct posix_acl * (*get_acl)(struct mnt_idmap *, struct dentry *,␣
↪int);
        int (*set_acl)(struct mnt_idmap *, struct dentry *, struct posix_acl *,
↪ int);
        int (*fileattr_set)(struct mnt_idmap *idmap,
                            struct dentry *dentry, struct fileattr *fa);
        int (*fileattr_get)(struct dentry *dentry, struct fileattr *fa);
        struct offset_ctx *(*get_offset_ctx)(struct inode *inode);
};
```

Again, all methods are called without any locks being held, unless otherwise noted.

**create**
: called by the open(2) and creat(2) system calls. Only required if you want to support regular files. The dentry you get should not have an inode (i.e. it should be a negative dentry). Here you will probably call *d_instantiate()* with the dentry and the newly created inode

**lookup**
: called when the VFS needs to look up an inode in a parent directory. The name to look for is found in the dentry. This method must call *d_add()* to insert the found inode into the dentry. The "i_count" field in the inode structure should be incremented. If the named inode does not exist a NULL inode should be inserted into the dentry (this is called a negative dentry). Returning an error code from this routine must only be done on a real error, otherwise creating inodes with system calls like create(2), mknod(2), mkdir(2) and so on will fail. If you wish to overload the dentry methods then you should initialise the "d_dop" field in the dentry; this is a pointer to a struct "dentry_operations". This method is called with the directory inode semaphore held

**link**
called by the link(2) system call. Only required if you want to support hard links. You will probably need to call *d_instantiate()* just as you would in the create() method

**unlink**
called by the unlink(2) system call. Only required if you want to support deleting inodes

**symlink**
called by the symlink(2) system call. Only required if you want to support symlinks. You will probably need to call *d_instantiate()* just as you would in the create() method

**mkdir**
called by the mkdir(2) system call. Only required if you want to support creating subdirectories. You will probably need to call *d_instantiate()* just as you would in the create() method

**rmdir**
called by the rmdir(2) system call. Only required if you want to support deleting subdirectories

**mknod**
called by the mknod(2) system call to create a device (char, block) inode or a named pipe (FIFO) or socket. Only required if you want to support creating these types of inodes. You will probably need to call *d_instantiate()* just as you would in the create() method

**rename**
called by the rename(2) system call to rename the object to have the parent and name given by the second inode and dentry.

The filesystem must return -EINVAL for any unsupported or unknown flags. Currently the following flags are implemented: (1) RENAME_NOREPLACE: this flag indicates that if the target of the rename exists the rename should fail with -EEXIST instead of replacing the target. The VFS already checks for existence, so for local filesystems the RENAME_NOREPLACE implementation is equivalent to plain rename. (2) RENAME_EXCHANGE: exchange source and target. Both must exist; this is checked by the VFS. Unlike plain rename, source and target may be of different type.

**get_link**
called by the VFS to follow a symbolic link to the inode it points to. Only required if you want to support symbolic links. This method returns the symlink body to traverse (and possibly resets the current position with nd_jump_link()). If the body won't go away until the inode is gone, nothing else is needed; if it needs to be otherwise pinned, arrange for its release by having get_link(..., ..., done) do set_delayed_call(done, destructor, argument). In that case destructor(argument) will be called once VFS is done with the body you've returned. May be called in RCU mode; that is indicated by NULL dentry argument. If request can't be handled without leaving RCU mode, have it return ERR_PTR(-ECHILD).

If the filesystem stores the symlink target in ->i_link, the VFS may use it directly without calling ->get_link(); however, ->get_link() must still be provided. ->i_link must not be freed until after an RCU grace period. Writing to ->i_link post-iget() time requires a 'release' memory barrier.

**readlink**
this is now just an override for use by readlink(2) for the cases when ->get_link uses nd_jump_link() or object is not in fact a symlink. Normally filesystems should only implement ->get_link for symlinks and readlink(2) will automatically use that.

**permission**
called by the VFS to check for access rights on a POSIX-like filesystem.

May be called in rcu-walk mode (mask & MAY_NOT_BLOCK). If in rcu-walk mode, the filesystem must check the permission without blocking or storing to the inode.

If a situation is encountered that rcu-walk cannot handle, return -ECHILD and it will be called again in ref-walk mode.

**setattr**
called by the VFS to set attributes for a file. This method is called by chmod(2) and related system calls.

**getattr**
called by the VFS to get attributes of a file. This method is called by stat(2) and related system calls.

**listxattr**
called by the VFS to list all extended attributes for a given file. This method is called by the listxattr(2) system call.

**update_time**
called by the VFS to update a specific time or the i_version of an inode. If this is not defined the VFS will update the inode itself and call mark_inode_dirty_sync.

**atomic_open**
called on the last component of an open. Using this optional method the filesystem can look up, possibly create and open the file in one atomic operation. If it wants to leave actual opening to the caller (e.g. if the file turned out to be a symlink, device, or just something filesystem won't do atomic open for), it may signal this by returning finish_no_open(file, dentry). This method is only called if the last component is negative or needs lookup. Cached positive dentries are still handled by f_op->open(). If the file was created, FMODE_CREATED flag should be set in file->f_mode. In case of O_EXCL the method must only succeed if the file didn't exist and hence FMODE_CREATED shall always be set on success.

**tmpfile**
called in the end of O_TMPFILE open(). Optional, equivalent to atomically creating, opening and unlinking a file in given directory. On success needs to return with the file already open; this can be done by calling finish_open_simple() right at the end.

**fileattr_get**
called on ioctl(FS_IOC_GETFLAGS) and ioctl(FS_IOC_FSGETXATTR) to retrieve miscellaneous file flags and attributes. Also called before the relevant SET operation to check what is being changed (in this case with i_rwsem locked exclusive). If unset, then fall back to f_op->ioctl().

**fileattr_set**
called on ioctl(FS_IOC_SETFLAGS) and ioctl(FS_IOC_FSSETXATTR) to change miscellaneous file flags and attributes. Callers hold i_rwsem exclusive. If unset, then fall back to f_op->ioctl().

**get_offset_ctx**
called to get the offset context for a directory inode. A filesystem must define this operation to use simple_offset_dir_operations.

### 1.1.5 The Address Space Object

The address space object is used to group and manage pages in the page cache. It can be used to keep track of the pages in a file (or anything else) and also track the mapping of sections of the file into process address spaces.

There are a number of distinct yet related services that an address-space can provide. These include communicating memory pressure, page lookup by address, and keeping track of pages tagged as Dirty or Writeback.

The first can be used independently to the others. The VM can try to either write dirty pages in order to clean them, or release clean pages in order to reuse them. To do this it can call the ->writepage method on dirty pages, and ->release_folio on clean folios with the private flag set. Clean pages without PagePrivate and with no external references will be released without notice being given to the address_space.

To achieve this functionality, pages need to be placed on an LRU with lru_cache_add and mark_page_active needs to be called whenever the page is used.

Pages are normally kept in a radix tree index by ->index. This tree maintains information about the PG_Dirty and PG_Writeback status of each page, so that pages with either of these flags can be found quickly.

The Dirty tag is primarily used by mpage_writepages - the default ->writepages method. It uses the tag to find dirty pages to call ->writepage on. If mpage_writepages is not used (i.e. the address provides its own ->writepages) , the PAGECACHE_TAG_DIRTY tag is almost unused. write_inode_now and sync_inode do use it (through __sync_single_inode) to check if ->writepages has been successful in writing out the whole address_space.

The Writeback tag is used by filemap*wait* and sync_page* functions, via filemap_fdatawait_range, to wait for all writeback to complete.

An address_space handler may attach extra information to a page, typically using the 'private' field in the 'struct page'. If such information is attached, the PG_Private flag should be set. This will cause various VM routines to make extra calls into the address_space handler to deal with that data.

An address space acts as an intermediate between storage and application. Data is read into the address space a whole page at a time, and provided to the application either by copying of the page, or by memory-mapping the page. Data is written into the address space by the application, and then written-back to storage typically in whole pages, however the address_space has finer control of write sizes.

The read process essentially only requires 'read_folio'. The write process is more complicated and uses write_begin/write_end or dirty_folio to write data into the address_space, and writepage and writepages to writeback data to storage.

Adding and removing pages to/from an address_space is protected by the inode's i_mutex.

When data is written to a page, the PG_Dirty flag should be set. It typically remains set until writepage asks for it to be written. This should clear PG_Dirty and set PG_Writeback. It can be actually written at any point after PG_Dirty is clear. Once it is known to be safe, PG_Writeback is cleared.

Writeback makes use of a writeback_control structure to direct the operations. This gives the writepage and writepages operations some information about the nature of and reason for the

writeback request, and the constraints under which it is being done. It is also used to return information back to the caller about the result of a writepage or writepages request.

### Handling errors during writeback

Most applications that do buffered I/O will periodically call a file synchronization call (fsync, fdatasync, msync or sync_file_range) to ensure that data written has made it to the backing store. When there is an error during writeback, they expect that error to be reported when a file sync request is made. After an error has been reported on one request, subsequent requests on the same file descriptor should return 0, unless further writeback errors have occurred since the previous file synchronization.

Ideally, the kernel would report errors only on file descriptions on which writes were done that subsequently failed to be written back. The generic pagecache infrastructure does not track the file descriptions that have dirtied each individual page however, so determining which file descriptors should get back an error is not possible.

Instead, the generic writeback error tracking infrastructure in the kernel settles for reporting errors to fsync on all file descriptions that were open at the time that the error occurred. In a situation with multiple writers, all of them will get back an error on a subsequent fsync, even if all of the writes done through that particular file descriptor succeeded (or even if there were no writes on that file descriptor at all).

Filesystems that wish to use this infrastructure should call mapping_set_error to record the error in the address_space when it occurs. Then, after writing back data from the pagecache in their file->fsync operation, they should call file_check_and_advance_wb_err to ensure that the struct file's error cursor has advanced to the correct point in the stream of errors emitted by the backing device(s).

### struct address_space_operations

This describes how the VFS can manipulate mapping of a file to page cache in your filesystem. The following members are defined:

```
struct address_space_operations {
        int (*writepage)(struct page *page, struct writeback_control *wbc);
        int (*read_folio)(struct file *, struct folio *);
        int (*writepages)(struct address_space *, struct writeback_control *);
        bool (*dirty_folio)(struct address_space *, struct folio *);
        void (*readahead)(struct readahead_control *);
        int (*write_begin)(struct file *, struct address_space *mapping,
                           loff_t pos, unsigned len,
                        struct page **pagep, void **fsdata);
        int (*write_end)(struct file *, struct address_space *mapping,
                         loff_t pos, unsigned len, unsigned copied,
                         struct page *page, void *fsdata);
        sector_t (*bmap)(struct address_space *, sector_t);
        void (*invalidate_folio) (struct folio *, size_t start, size_t len);
        bool (*release_folio)(struct folio *, gfp_t);
        void (*free_folio)(struct folio *);
        ssize_t (*direct_IO)(struct kiocb *, struct iov_iter *iter);
        int (*migrate_folio)(struct mapping *, struct folio *dst,
```

```
                        struct folio *src, enum migrate_mode);
        int (*launder_folio) (struct folio *);

        bool (*is_partially_uptodate) (struct folio *, size_t from,
                                       size_t count);
        void (*is_dirty_writeback)(struct folio *, bool *, bool *);
        int (*error_remove_folio)(struct mapping *mapping, struct folio *);
        int (*swap_activate)(struct swap_info_struct *sis, struct file *f,␣
↪sector_t *span)
        int (*swap_deactivate)(struct file *);
        int (*swap_rw)(struct kiocb *iocb, struct iov_iter *iter);
};
```

**writepage**
called by the VM to write a dirty page to backing store. This may happen for data integrity reasons (i.e. 'sync'), or to free up memory (flush). The difference can be seen in wbc->sync_mode. The PG_Dirty flag has been cleared and PageLocked is true. writepage should start writeout, should set PG_Writeback, and should make sure the page is unlocked, either synchronously or asynchronously when the write operation completes.

If wbc->sync_mode is WB_SYNC_NONE, ->writepage doesn't have to try too hard if there are problems, and may choose to write out other pages from the mapping if that is easier (e.g. due to internal dependencies). If it chooses not to start writeout, it should return AOP_WRITEPAGE_ACTIVATE so that the VM will not keep calling ->writepage on that page.

See the file "Locking" for more details.

**read_folio**
Called by the page cache to read a folio from the backing store. The 'file' argument supplies authentication information to network filesystems, and is generally not used by block based filesystems. It may be NULL if the caller does not have an open file (eg if the kernel is performing a read for itself rather than on behalf of a userspace process with an open file).

If the mapping does not support large folios, the folio will contain a single page. The folio will be locked when read_folio is called. If the read completes successfully, the folio should be marked uptodate. The filesystem should unlock the folio once the read has completed, whether it was successful or not. The filesystem does not need to modify the refcount on the folio; the page cache holds a reference count and that will not be released until the folio is unlocked.

Filesystems may implement ->read_folio() synchronously. In normal operation, folios are read through the ->readahead() method. Only if this fails, or if the caller needs to wait for the read to complete will the page cache call ->read_folio(). Filesystems should not attempt to perform their own readahead in the ->read_folio() operation.

If the filesystem cannot perform the read at this time, it can unlock the folio, do whatever action it needs to ensure that the read will succeed in the future and return AOP_TRUNCATED_PAGE. In this case, the caller should look up the folio, lock it, and call ->read_folio again.

Callers may invoke the ->read_folio() method directly, but using read_mapping_folio() will take care of locking, waiting for the read to complete and handle cases such as

AOP_TRUNCATED_PAGE.

**writepages**
called by the VM to write out pages associated with the address_space object. If wbc->sync_mode is WB_SYNC_ALL, then the writeback_control will specify a range of pages that must be written out. If it is WB_SYNC_NONE, then a nr_to_write is given and that many pages should be written if possible. If no ->writepages is given, then mpage_writepages is used instead. This will choose pages from the address space that are tagged as DIRTY and will pass them to ->writepage.

**dirty_folio**
called by the VM to mark a folio as dirty. This is particularly needed if an address space attaches private data to a folio, and that data needs to be updated when a folio is dirtied. This is called, for example, when a memory mapped page gets modified. If defined, it should set the folio dirty flag, and the PAGECACHE_TAG_DIRTY search mark in i_pages.

**readahead**
Called by the VM to read pages associated with the address_space object. The pages are consecutive in the page cache and are locked. The implementation should decrement the page refcount after starting I/O on each page. Usually the page will be unlocked by the I/O completion handler. The set of pages are divided into some sync pages followed by some async pages, rac->ra->async_size gives the number of async pages. The filesystem should attempt to read all sync pages but may decide to stop once it reaches the async pages. If it does decide to stop attempting I/O, it can simply return. The caller will remove the remaining pages from the address space, unlock them and decrement the page refcount. Set PageUptodate if the I/O completes successfully. Setting PageError on any page will be ignored; simply unlock the page if an I/O error occurs.

**write_begin**
Called by the generic buffered write code to ask the filesystem to prepare to write len bytes at the given offset in the file. The address_space should check that the write will be able to complete, by allocating space if necessary and doing any other internal housekeeping. If the write will update parts of any basic-blocks on storage, then those blocks should be pre-read (if they haven't been read already) so that the updated blocks can be written out properly.

The filesystem must return the locked pagecache page for the specified offset, in `*pagep`, for the caller to write into.

It must be able to cope with short writes (where the length passed to write_begin is greater than the number of bytes copied into the page).

A void * may be returned in fsdata, which then gets passed into write_end.

Returns 0 on success; < 0 on failure (which is the error code), in which case write_end is not called.

**write_end**
After a successful write_begin, and data copy, write_end must be called. len is the original len passed to write_begin, and copied is the amount that was able to be copied.

The filesystem must take care of unlocking the page and releasing it refcount, and updating i_size.

Returns < 0 on failure, otherwise the number of bytes (<= 'copied') that were able to be copied into pagecache.

**bmap**
called by the VFS to map a logical block offset within object to physical block number. This method is used by the FIBMAP ioctl and for working with swap-files. To be able to swap to a file, the file must have a stable mapping to a block device. The swap system does not go through the filesystem but instead uses bmap to find out where the blocks in the file are and uses those addresses directly.

**invalidate_folio**
If a folio has private data, then invalidate_folio will be called when part or all of the folio is to be removed from the address space. This generally corresponds to either a truncation, punch hole or a complete invalidation of the address space (in the latter case 'offset' will always be 0 and 'length' will be folio_size()). Any private data associated with the folio should be updated to reflect this truncation. If offset is 0 and length is folio_size(), then the private data should be released, because the folio must be able to be completely discarded. This may be done by calling the ->release_folio function, but in this case the release MUST succeed.

**release_folio**
release_folio is called on folios with private data to tell the filesystem that the folio is about to be freed. ->release_folio should remove any private data from the folio and clear the private flag. If release_folio() fails, it should return false. release_folio() is used in two distinct though related cases. The first is when the VM wants to free a clean folio with no active users. If ->release_folio succeeds, the folio will be removed from the address_space and be freed.

The second case is when a request has been made to invalidate some or all folios in an address_space. This can happen through the fadvise(POSIX_FADV_DONTNEED) system call or by the filesystem explicitly requesting it as nfs and 9p do (when they believe the cache may be out of date with storage) by calling invalidate_inode_pages2(). If the filesystem makes such a call, and needs to be certain that all folios are invalidated, then its release_folio will need to ensure this. Possibly it can clear the uptodate flag if it cannot free private data yet.

**free_folio**
free_folio is called once the folio is no longer visible in the page cache in order to allow the cleanup of any private data. Since it may be called by the memory reclaimer, it should not assume that the original address_space mapping still exists, and it should not block.

**direct_IO**
called by the generic read/write routines to perform direct_IO - that is IO requests which bypass the page cache and transfer data directly between the storage and the application's address space.

**migrate_folio**
This is used to compact the physical memory usage. If the VM wants to relocate a folio (maybe from a memory device that is signalling imminent failure) it will pass a new folio and an old folio to this function. migrate_folio should transfer any private data across and update any references that it has to the folio.

**launder_folio**
Called before freeing a folio - it writes back the dirty folio. To prevent redirtying the folio, it is kept locked during the whole operation.

**is_partially_uptodate**
Called by the VM when reading a file through the pagecache when the underlying blocksize

is smaller than the size of the folio. If the required block is up to date then the read can complete without needing I/O to bring the whole page up to date.

**is_dirty_writeback**
Called by the VM when attempting to reclaim a folio. The VM uses dirty and writeback information to determine if it needs to stall to allow flushers a chance to complete some IO. Ordinarily it can use folio_test_dirty and folio_test_writeback but some filesystems have more complex state (unstable folios in NFS prevent reclaim) or do not set those flags due to locking problems. This callback allows a filesystem to indicate to the VM if a folio should be treated as dirty or writeback for the purposes of stalling.

**error_remove_folio**
normally set to generic_error_remove_folio if truncation is ok for this address space. Used for memory failure handling. Setting this implies you deal with pages going away under you, unless you have them locked or reference counts increased.

`swap_activate`

Called to prepare the given file for swap. It should perform any validation and preparation necessary to ensure that writes can be performed with minimal memory allocation. It should call add_swap_extent(), or the helper iomap_swapfile_activate(), and return the number of extents added. If IO should be submitted through ->swap_rw(), it should set SWP_FS_OPS, otherwise IO will be submitted directly to the block device `sis->bdev`.

**swap_deactivate**
Called during swapoff on files where swap_activate was successful.

**swap_rw**
Called to read or write swap pages when SWP_FS_OPS is set.

### 1.1.6 The File Object

A file object represents a file opened by a process. This is also known as an "open file description" in POSIX parlance.

#### struct file_operations

This describes how the VFS can manipulate an open file. As of kernel 4.18, the following members are defined:

```
struct file_operations {
        struct module *owner;
        loff_t (*llseek) (struct file *, loff_t, int);
        ssize_t (*read) (struct file *, char __user *, size_t, loff_t *);
        ssize_t (*write) (struct file *, const char __user *, size_t, loff_t␣
↪*);
        ssize_t (*read_iter) (struct kiocb *, struct iov_iter *);
        ssize_t (*write_iter) (struct kiocb *, struct iov_iter *);
        int (*iopoll)(struct kiocb *kiocb, bool spin);
        int (*iterate_shared) (struct file *, struct dir_context *);
        __poll_t (*poll) (struct file *, struct poll_table_struct *);
        long (*unlocked_ioctl) (struct file *, unsigned int, unsigned long);
```

```
        long (*compat_ioctl) (struct file *, unsigned int, unsigned long);
        int (*mmap) (struct file *, struct vm_area_struct *);
        int (*open) (struct inode *, struct file *);
        int (*flush) (struct file *, fl_owner_t id);
        int (*release) (struct inode *, struct file *);
        int (*fsync) (struct file *, loff_t, loff_t, int datasync);
        int (*fasync) (int, struct file *, int);
        int (*lock) (struct file *, int, struct file_lock *);
        unsigned long (*get_unmapped_area)(struct file *, unsigned long,␣
↪unsigned long, unsigned long, unsigned long);
        int (*check_flags)(int);
        int (*flock) (struct file *, int, struct file_lock *);
        ssize_t (*splice_write)(struct pipe_inode_info *, struct file *, loff_
↪t *, size_t, unsigned int);
        ssize_t (*splice_read)(struct file *, loff_t *, struct pipe_inode_info␣
↪*, size_t, unsigned int);
        int (*setlease)(struct file *, long, struct file_lock **, void **);
        long (*fallocate)(struct file *file, int mode, loff_t offset,
                          loff_t len);
        void (*show_fdinfo)(struct seq_file *m, struct file *f);
#ifndef CONFIG_MMU
        unsigned (*mmap_capabilities)(struct file *);
#endif
        ssize_t (*copy_file_range)(struct file *, loff_t, struct file *, loff_
↪t, size_t, unsigned int);
        loff_t (*remap_file_range)(struct file *file_in, loff_t pos_in,
                                   struct file *file_out, loff_t pos_out,
                                   loff_t len, unsigned int remap_flags);
        int (*fadvise)(struct file *, loff_t, loff_t, int);
};
```

Again, all methods are called without any locks being held, unless otherwise noted.

**llseek**
called when the VFS needs to move the file position index

**read**
called by read(2) and related system calls

**read_iter**
possibly asynchronous read with iov_iter as destination

**write**
called by write(2) and related system calls

**write_iter**
possibly asynchronous write with iov_iter as source

**iopoll**
called when aio wants to poll for completions on HIPRI iocbs

**iterate_shared**
called when the VFS needs to read the directory contents

**poll**

called by the VFS when a process wants to check if there is activity on this file and (optionally) go to sleep until there is activity. Called by the select(2) and poll(2) system calls

**unlocked_ioctl**
called by the ioctl(2) system call.

**compat_ioctl**

**called by the ioctl(2) system call when 32 bit system calls are**
used on 64 bit kernels.

**mmap**
called by the mmap(2) system call

**open**
called by the VFS when an inode should be opened. When the VFS opens a file, it creates a new ”struct file”. It then calls the open method for the newly allocated file structure. You might think that the open method really belongs in ”struct inode_operations”, and you may be right. I think it's done the way it is because it makes filesystems simpler to implement. The open() method is a good place to initialize the ”private_data” member in the file structure if you want to point to a device structure

**flush**
called by the close(2) system call to flush a file

**release**
called when the last reference to an open file is closed

**fsync**
called by the fsync(2) system call. Also see the section above entitled ”Handling errors during writeback”.

**fasync**
called by the fcntl(2) system call when asynchronous (non-blocking) mode is enabled for a file

**lock**
called by the fcntl(2) system call for F_GETLK, F_SETLK, and F_SETLKW commands

**get_unmapped_area**
called by the mmap(2) system call

**check_flags**
called by the fcntl(2) system call for F_SETFL command

**flock**
called by the flock(2) system call

**splice_write**
called by the VFS to splice data from a pipe to a file. This method is used by the splice(2) system call

**splice_read**
called by the VFS to splice data from file to a pipe. This method is used by the splice(2) system call

**setlease**
called by the VFS to set or release a file lock lease. setlease implementations should call generic_setlease to record or remove the lease in the inode after setting it.

**fallocate**
called by the VFS to preallocate blocks or punch a hole.

**copy_file_range**
called by the copy_file_range(2) system call.

**remap_file_range**
called by the ioctl(2) system call for FICLONERANGE and FICLONE and FIDEDUPERANGE commands to remap file ranges. An implementation should remap len bytes at pos_in of the source file into the dest file at pos_out. Implementations must handle callers passing in len == 0; this means "remap to the end of the source file". The return value should the number of bytes remapped, or the usual negative error code if errors occurred before any bytes were remapped. The remap_flags parameter accepts REMAP_FILE_* flags. If REMAP_FILE_DEDUP is set then the implementation must only remap if the requested file ranges have identical contents. If REMAP_FILE_CAN_SHORTEN is set, the caller is ok with the implementation shortening the request length to satisfy alignment or EOF requirements (or any other reason).

**fadvise**
possibly called by the fadvise64() system call.

Note that the file operations are implemented by the specific filesystem in which the inode resides. When opening a device node (character or block special) most filesystems will call special support routines in the VFS which will locate the required device driver information. These support routines replace the filesystem file operations with those for the device driver, and then proceed to call the new open() method for the file. This is how opening a device file in the filesystem eventually ends up calling the device driver open() method.

### 1.1.7 Directory Entry Cache (dcache)

#### struct dentry_operations

This describes how a filesystem can overload the standard dentry operations. Dentries and the dcache are the domain of the VFS and the individual filesystem implementations. Device drivers have no business here. These methods may be set to NULL, as they are either optional or the VFS uses a default. As of kernel 2.6.22, the following members are defined:

```
struct dentry_operations {
        int (*d_revalidate)(struct dentry *, unsigned int);
        int (*d_weak_revalidate)(struct dentry *, unsigned int);
        int (*d_hash)(const struct dentry *, struct qstr *);
        int (*d_compare)(const struct dentry *,
                         unsigned int, const char *, const struct qstr *);
        int (*d_delete)(const struct dentry *);
        int (*d_init)(struct dentry *);
        void (*d_release)(struct dentry *);
        void (*d_iput)(struct dentry *, struct inode *);
        char *(*d_dname)(struct dentry *, char *, int);
        struct vfsmount *(*d_automount)(struct path *);
        int (*d_manage)(const struct path *, bool);
        struct dentry *(*d_real)(struct dentry *, const struct inode *);
};
```

**d_revalidate**
called when the VFS needs to revalidate a dentry. This is called whenever a name look-up finds a dentry in the dcache. Most local filesystems leave this as NULL, because all their dentries in the dcache are valid. Network filesystems are different since things can change on the server without the client necessarily being aware of it.

This function should return a positive value if the dentry is still valid, and zero or a negative error code if it isn't.

d_revalidate may be called in rcu-walk mode (flags & LOOKUP_RCU). If in rcu-walk mode, the filesystem must revalidate the dentry without blocking or storing to the dentry, d_parent and d_inode should not be used without care (because they can change and, in d_inode case, even become NULL under us).

If a situation is encountered that rcu-walk cannot handle, return -ECHILD and it will be called again in ref-walk mode.

**d_weak_revalidate**
called when the VFS needs to revalidate a "jumped" dentry. This is called when a path-walk ends at dentry that was not acquired by doing a lookup in the parent directory. This includes "/", "." and "..", as well as procfs-style symlinks and mountpoint traversal.

In this case, we are less concerned with whether the dentry is still fully correct, but rather that the inode is still valid. As with d_revalidate, most local filesystems will set this to NULL since their dcache entries are always valid.

This function has the same return code semantics as d_revalidate.

d_weak_revalidate is only called after leaving rcu-walk mode.

**d_hash**
called when the VFS adds a dentry to the hash table. The first dentry passed to d_hash is the parent directory that the name is to be hashed into.

Same locking and synchronisation rules as d_compare regarding what is safe to dereference etc.

**d_compare**
called to compare a dentry name with a given name. The first dentry is the parent of the dentry to be compared, the second is the child dentry. len and name string are properties of the dentry to be compared. qstr is the name to compare it with.

Must be constant and idempotent, and should not take locks if possible, and should not or store into the dentry. Should not dereference pointers outside the dentry without lots of care (eg. d_parent, d_inode, d_name should not be used).

However, our vfsmount is pinned, and RCU held, so the dentries and inodes won't disappear, neither will our sb or filesystem module. ->d_sb may be used.

It is a tricky calling convention because it needs to be called under "rcu-walk", ie. without any locks or references on things.

**d_delete**
called when the last reference to a dentry is dropped and the dcache is deciding whether or not to cache it. Return 1 to delete immediately, or 0 to cache the dentry. Default is NULL which means to always cache a reachable dentry. d_delete must be constant and idempotent.

**d_init**
called when a dentry is allocated

**d_release**
called when a dentry is really deallocated

**d_iput**
called when a dentry loses its inode (just prior to its being deallocated). The default when this is NULL is that the VFS calls *iput()*. If you define this method, you must call *iput()* yourself

**d_dname**
called when the pathname of a dentry should be generated. Useful for some pseudo filesystems (sockfs, pipefs, ...) to delay pathname generation. (Instead of doing it when dentry is created, it's done only when the path is needed.). Real filesystems probably dont want to use it, because their dentries are present in global dcache hash, so their hash should be an invariant. As no lock is held, d_dname() should not try to modify the dentry itself, unless appropriate SMP safety is used. CAUTION : *d_path()* logic is quite tricky. The correct way to return for example "Hello" is to put it at the end of the buffer, and returns a pointer to the first char. dynamic_dname() helper function is provided to take care of this.

Example :

```
static char *pipefs_dname(struct dentry *dent, char *buffer, int buflen)
{
        return dynamic_dname(dentry, buffer, buflen, "pipe:[%lu]",
                             dentry->d_inode->i_ino);
}
```

**d_automount**
called when an automount dentry is to be traversed (optional). This should create a new VFS mount record and return the record to the caller. The caller is supplied with a path parameter giving the automount directory to describe the automount target and the parent VFS mount record to provide inheritable mount parameters. NULL should be returned if someone else managed to make the automount first. If the vfsmount creation failed, then an error code should be returned. If -EISDIR is returned, then the directory will be treated as an ordinary directory and returned to pathwalk to continue walking.

If a vfsmount is returned, the caller will attempt to mount it on the mountpoint and will remove the vfsmount from its expiration list in the case of failure. The vfsmount should be returned with 2 refs on it to prevent automatic expiration - the caller will clean up the additional ref.

This function is only used if DCACHE_NEED_AUTOMOUNT is set on the dentry. This is set by __d_instantiate() if S_AUTOMOUNT is set on the inode being added.

**d_manage**
called to allow the filesystem to manage the transition from a dentry (optional). This allows autofs, for example, to hold up clients waiting to explore behind a 'mountpoint' while letting the daemon go past and construct the subtree there. 0 should be returned to let the calling process continue. -EISDIR can be returned to tell pathwalk to use this directory as an ordinary directory and to ignore anything mounted on it and not to check the automount flag. Any other error code will abort pathwalk completely.

If the 'rcu_walk' parameter is true, then the caller is doing a pathwalk in RCU-walk mode.

Sleeping is not permitted in this mode, and the caller can be asked to leave it and call again by returning -ECHILD. -EISDIR may also be returned to tell pathwalk to ignore d_automount or any mounts.

This function is only used if DCACHE_MANAGE_TRANSIT is set on the dentry being transited from.

**d_real**
overlay/union type filesystems implement this method to return one of the underlying dentries hidden by the overlay. It is used in two different modes:

Called from file_dentry() it returns the real dentry matching the inode argument. The real dentry may be from a lower layer already copied up, but still referenced from the file. This mode is selected with a non-NULL inode argument.

With NULL inode the topmost real underlying dentry is returned.

Each dentry has a pointer to its parent dentry, as well as a hash list of child dentries. Child dentries are basically like files in a directory.

### Directory Entry Cache API

There are a number of functions defined which permit a filesystem to manipulate dentries:

**dget**
open a new handle for an existing dentry (this just increments the usage count)

**dput**
close a handle for a dentry (decrements the usage count). If the usage count drops to 0, and the dentry is still in its parent's hash, the ”d_delete” method is called to check whether it should be cached. If it should not be cached, or if the dentry is not hashed, it is deleted. Otherwise cached dentries are put into an LRU list to be reclaimed on memory shortage.

**d_drop**
this unhashes a dentry from its parents hash list. A subsequent call to dput() will deallocate the dentry if its usage count drops to 0

**d_delete**
delete a dentry. If there are no other open references to the dentry then the dentry is turned into a negative dentry (the d_iput() method is called). If there are other references, then *d_drop()* is called instead

**d_add**
add a dentry to its parents hash list and then calls *d_instantiate()*

**d_instantiate**
add a dentry to the alias hash list for the inode and updates the ”d_inode” member. The ”i_count” member in the inode structure should be set/incremented. If the inode pointer is NULL, the dentry is called a ”negative dentry”. This function is commonly called when an inode is created for an existing negative dentry

**d_lookup**
look up a dentry given its parent and path name component It looks up the child of that given name from the dcache hash table. If it is found, the reference count is incremented and the dentry is returned. The caller must use dput() to free the dentry when it finishes using it.

### 1.1.8 Mount Options

#### Parsing options

On mount and remount the filesystem is passed a string containing a comma separated list of mount options. The options can have either of these forms:

    option option=value

The <linux/parser.h> header defines an API that helps parse these options. There are plenty of examples on how to use it in existing filesystems.

#### Showing options

If a filesystem accepts mount options, it must define show_options() to show all the currently active options. The rules are:

- options MUST be shown which are not default or their values differ from the default
- options MAY be shown which are enabled by default or have their default value

Options used only internally between a mount helper and the kernel (such as file descriptors), or which only have an effect during the mounting (such as ones controlling the creation of a journal) are exempt from the above rules.

The underlying reason for the above rules is to make sure, that a mount can be accurately replicated (e.g. umounting and mounting again) based on the information found in /proc/mounts.

### 1.1.9 Resources

**(Note some of these resources are not up-to-date with the latest kernel**
    version.)

**Creating Linux virtual filesystems. 2002**
    <https://lwn.net/Articles/13325/>

**The Linux Virtual File-system Layer by Neil Brown. 1999**
    <http://www.cse.unsw.edu.au/~neilb/oss/linux-commentary/vfs.html>

**A tour of the Linux VFS by Michael K. Johnson. 1996**
    <https://www.tldp.org/LDP/khg/HyperNews/get/fs/vfstour.html>

**A small trail through the Linux kernel by Andries Brouwer. 2001**
    <https://www.win.tue.nl/~aeb/linux/vfs/trail.html>

## 1.2 Pathname lookup

This write-up is based on three articles published at lwn.net:

- <https://lwn.net/Articles/649115/> Pathname lookup in Linux
- <https://lwn.net/Articles/649729/> RCU-walk: faster pathname lookup in Linux
- <https://lwn.net/Articles/650786/> A walk among the symlinks

Written by Neil Brown with help from Al Viro and Jon Corbet. It has subsequently been updated to reflect changes in the kernel including:

- per-directory parallel name lookup.
- `openat2()` resolution restriction flags.

### 1.2.1 Introduction to pathname lookup

The most obvious aspect of pathname lookup, which very little exploration is needed to discover, is that it is complex. There are many rules, special cases, and implementation alternatives that all combine to confuse the unwary reader. Computer science has long been acquainted with such complexity and has tools to help manage it. One tool that we will make extensive use of is ”divide and conquer”. For the early parts of the analysis we will divide off symlinks - leaving them until the final part. Well before we get to symlinks we have another major division based on the VFS's approach to locking which will allow us to review ”REF-walk” and ”RCU-walk” separately. But we are getting ahead of ourselves. There are some important low level distinctions we need to clarify first.

#### There are two sorts of ...

Pathnames (sometimes ”file names”), used to identify objects in the filesystem, will be familiar to most readers. They contain two sorts of elements: ”slashes” that are sequences of one or more ”`/`” characters, and ”components” that are sequences of one or more non-”`/`” characters. These form two kinds of paths. Those that start with slashes are ”absolute” and start from the filesystem root. The others are ”relative” and start from the current directory, or from some other location specified by a file descriptor given to ”`*at()`” system calls such as openat().

It is tempting to describe the second kind as starting with a component, but that isn't always accurate: a pathname can lack both slashes and components, it can be empty, in other words. This is generally forbidden in POSIX, but some of those ”`*at()`” system calls in Linux permit it when the `AT_EMPTY_PATH` flag is given. For example, if you have an open file descriptor on an executable file you can execute it by calling execveat() passing the file descriptor, an empty path, and the `AT_EMPTY_PATH` flag.

These paths can be divided into two sections: the final component and everything else. The ”everything else” is the easy bit. In all cases it must identify a directory that already exists, otherwise an error such as `ENOENT` or `ENOTDIR` will be reported.

The final component is not so simple. Not only do different system calls interpret it quite differently (e.g. some create it, some do not), but it might not even exist: neither the empty pathname nor the pathname that is just slashes have a final component. If it does exist, it could be ”`.`” or ”`..`” which are handled quite differently from other components.

If a pathname ends with a slash, such as "/tmp/foo/" it might be tempting to consider that to have an empty final component. In many ways that would lead to correct results, but not always. In particular, mkdir() and rmdir() each create or remove a directory named by the final component, and they are required to work with pathnames ending in "/". According to POSIX:

> A pathname that contains at least one non-<slash> character and that ends with one or more trailing <slash> characters shall not be resolved successfully unless the last pathname component before the trailing <slash> characters names an existing directory or a directory entry that is to be created for a directory immediately after the pathname is resolved.

The Linux pathname walking code (mostly in fs/namei.c) deals with all of these issues: breaking the path into components, handling the "everything else" quite separately from the final component, and checking that the trailing slash is not used where it isn't permitted. It also addresses the important issue of concurrent access.

While one process is looking up a pathname, another might be making changes that affect that lookup. One fairly extreme case is that if "a/b" were renamed to "a/c/b" while another process were looking up "a/b/..", that process might successfully resolve on "a/c". Most races are much more subtle, and a big part of the task of pathname lookup is to prevent them from having damaging effects. Many of the possible races are seen most clearly in the context of the "dcache" and an understanding of that is central to understanding pathname lookup.

### More than just a cache

The "dcache" caches information about names in each filesystem to make them quickly available for lookup. Each entry (known as a "dentry") contains three significant fields: a component name, a pointer to a parent dentry, and a pointer to the "inode" which contains further information about the object in that parent with the given name. The inode pointer can be NULL indicating that the name doesn't exist in the parent. While there can be linkage in the dentry of a directory to the dentries of the children, that linkage is not used for pathname lookup, and so will not be considered here.

The dcache has a number of uses apart from accelerating lookup. One that will be particularly relevant is that it is closely integrated with the mount table that records which filesystem is mounted where. What the mount table actually stores is which dentry is mounted on top of which other dentry.

When considering the dcache, we have another of our "two types" distinctions: there are two types of filesystems.

Some filesystems ensure that the information in the dcache is always completely accurate (though not necessarily complete). This can allow the VFS to determine if a particular file does or doesn't exist without checking with the filesystem, and means that the VFS can protect the filesystem against certain races and other problems. These are typically "local" filesystems such as ext3, XFS, and Btrfs.

Other filesystems don't provide that guarantee because they cannot. These are typically filesystems that are shared across a network, whether remote filesystems like NFS and 9P, or cluster filesystems like ocfs2 or cephfs. These filesystems allow the VFS to revalidate cached information, and must provide their own protection against awkward races. The VFS can detect these filesystems by the DCACHE_OP_REVALIDATE flag being set in the dentry.

### REF-walk: simple concurrency management with refcounts and spinlocks

With all of those divisions carefully classified, we can now start looking at the actual process of walking along a path. In particular we will start with the handling of the ”everything else” part of a pathname, and focus on the ”REF-walk” approach to concurrency management. This code is found in the `link_path_walk()` function, if you ignore all the places that only run when ”`LOOKUP_RCU`” (indicating the use of RCU-walk) is set.

REF-walk is fairly heavy-handed with locks and reference counts. Not as heavy-handed as in the old ”big kernel lock” days, but certainly not afraid of taking a lock when one is needed. It uses a variety of different concurrency controls. A background understanding of the various primitives is assumed, or can be gleaned from elsewhere such as in Meet the Lockers.

The locking mechanisms used by REF-walk include:

#### dentry->d_lockref

This uses the lockref primitive to provide both a spinlock and a reference count. The special-sauce of this primitive is that the conceptual sequence ”lock; inc_ref; unlock;” can often be performed with a single atomic memory operation.

Holding a reference on a dentry ensures that the dentry won't suddenly be freed and used for something else, so the values in various fields will behave as expected. It also protects the `->d_inode` reference to the inode to some extent.

The association between a dentry and its inode is fairly permanent. For example, when a file is renamed, the dentry and inode move together to the new location. When a file is created the dentry will initially be negative (i.e. `d_inode` is `NULL`), and will be assigned to the new inode as part of the act of creation.

When a file is deleted, this can be reflected in the cache either by setting `d_inode` to `NULL`, or by removing it from the hash table (described shortly) used to look up the name in the parent directory. If the dentry is still in use the second option is used as it is perfectly legal to keep using an open file after it has been deleted and having the dentry around helps. If the dentry is not otherwise in use (i.e. if the refcount in `d_lockref` is one), only then will `d_inode` be set to `NULL`. Doing it this way is more efficient for a very common case.

So as long as a counted reference is held to a dentry, a non-`NULL` `->d_inode` value will never be changed.

#### dentry->d_lock

`d_lock` is a synonym for the spinlock that is part of `d_lockref` above. For our purposes, holding this lock protects against the dentry being renamed or unlinked. In particular, its parent (`d_parent`), and its name (`d_name`) cannot be changed, and it cannot be removed from the dentry hash table.

When looking for a name in a directory, REF-walk takes `d_lock` on each candidate dentry that it finds in the hash table and then checks that the parent and name are correct. So it doesn't lock the parent while searching in the cache; it only locks children.

When looking for the parent for a given name (to handle ”`..`”), REF-walk can take `d_lock` to get a stable reference to `d_parent`, but it first tries a more lightweight approach. As seen in

dget_parent(), if a reference can be claimed on the parent, and if subsequently d_parent can be seen to have not changed, then there is no need to actually take the lock on the child.

### rename_lock

Looking up a given name in a given directory involves computing a hash from the two values (the name and the dentry of the directory), accessing that slot in a hash table, and searching the linked list that is found there.

When a dentry is renamed, the name and the parent dentry can both change so the hash will almost certainly change too. This would move the dentry to a different chain in the hash table. If a filename search happened to be looking at a dentry that was moved in this way, it might end up continuing the search down the wrong chain, and so miss out on part of the correct chain.

The name-lookup process (d_lookup()) does *not* try to prevent this from happening, but only to detect when it happens. rename_lock is a seqlock that is updated whenever any dentry is renamed. If d_lookup finds that a rename happened while it unsuccessfully scanned a chain in the hash table, it simply tries again.

rename_lock is also used to detect and defend against potential attacks against LOOKUP_BENEATH and LOOKUP_IN_ROOT when resolving ".." (where the parent directory is moved outside the root, bypassing the path_equal() check). If rename_lock is updated during the lookup and the path encounters a "..", a potential attack occurred and handle_dots() will bail out with -EAGAIN.

### inode->i_rwsem

i_rwsem is a read/write semaphore that serializes all changes to a particular directory. This ensures that, for example, an unlink() and a rename() cannot both happen at the same time. It also keeps the directory stable while the filesystem is asked to look up a name that is not currently in the dcache or, optionally, when the list of entries in a directory is being retrieved with readdir().

This has a complementary role to that of d_lock: i_rwsem on a directory protects all of the names in that directory, while d_lock on a name protects just one name in a directory. Most changes to the dcache hold i_rwsem on the relevant directory inode and briefly take d_lock on one or more the dentries while the change happens. One exception is when idle dentries are removed from the dcache due to memory pressure. This uses d_lock, but i_rwsem plays no role.

The semaphore affects pathname lookup in two distinct ways. Firstly it prevents changes during lookup of a name in a directory. walk_component() uses lookup_fast() first which, in turn, checks to see if the name is in the cache, using only d_lock locking. If the name isn't found, then walk_component() falls back to lookup_slow() which takes a shared lock on i_rwsem, checks again that the name isn't in the cache, and then calls in to the filesystem to get a definitive answer. A new dentry will be added to the cache regardless of the result.

Secondly, when pathname lookup reaches the final component, it will sometimes need to take an exclusive lock on i_rwsem before performing the last lookup so that the required exclusion can be achieved. How path lookup chooses to take, or not take, i_rwsem is one of the issues addressed in a subsequent section.

If two threads attempt to look up the same name at the same time - a name that is not yet in the dcache - the shared lock on `i_rwsem` will not prevent them both adding new dentries with the same name. As this would result in confusion an extra level of interlocking is used, based around a secondary hash table (`in_lookup_hashtable`) and a per-dentry flag bit (`DCACHE_PAR_LOOKUP`).

To add a new dentry to the cache while only holding a shared lock on `i_rwsem`, a thread must call `d_alloc_parallel()`. This allocates a dentry, stores the required name and parent in it, checks if there is already a matching dentry in the primary or secondary hash tables, and if not, stores the newly allocated dentry in the secondary hash table, with `DCACHE_PAR_LOOKUP` set.

If a matching dentry was found in the primary hash table then that is returned and the caller can know that it lost a race with some other thread adding the entry. If no matching dentry is found in either cache, the newly allocated dentry is returned and the caller can detect this from the presence of `DCACHE_PAR_LOOKUP`. In this case it knows that it has won any race and now is responsible for asking the filesystem to perform the lookup and find the matching inode. When the lookup is complete, it must call `d_lookup_done()` which clears the flag and does some other house keeping, including removing the dentry from the secondary hash table - it will normally have been added to the primary hash table already. Note that a `struct waitqueue_head` is passed to `d_alloc_parallel()`, and `d_lookup_done()` must be called while this `waitqueue_head` is still in scope.

If a matching dentry is found in the secondary hash table, `d_alloc_parallel()` has a little more work to do. It first waits for `DCACHE_PAR_LOOKUP` to be cleared, using a wait_queue that was passed to the instance of `d_alloc_parallel()` that won the race and that will be woken by the call to `d_lookup_done()`. It then checks to see if the dentry has now been added to the primary hash table. If it has, the dentry is returned and the caller just sees that it lost any race. If it hasn't been added to the primary hash table, the most likely explanation is that some other dentry was added instead using `d_splice_alias()`. In any case, `d_alloc_parallel()` repeats all the look ups from the start and will normally return something from the primary hash table.

### mnt->mnt_count

`mnt_count` is a per-CPU reference counter on "`mount`" structures. Per-CPU here means that incrementing the count is cheap as it only uses CPU-local memory, but checking if the count is zero is expensive as it needs to check with every CPU. Taking a `mnt_count` reference prevents the mount structure from disappearing as the result of regular unmount operations, but does not prevent a "lazy" unmount. So holding `mnt_count` doesn't ensure that the mount remains in the namespace and, in particular, doesn't stabilize the link to the mounted-on dentry. It does, however, ensure that the `mount` data structure remains coherent, and it provides a reference to the root dentry of the mounted filesystem. So a reference through `->mnt_count` provides a stable reference to the mounted dentry, but not the mounted-on dentry.

#### mount_lock

`mount_lock` is a global seqlock, a bit like `rename_lock`. It can be used to check if any change has been made to any mount points.

While walking down the tree (away from the root) this lock is used when crossing a mount point to check that the crossing was safe. That is, the value in the seqlock is read, then the code finds the mount that is mounted on the current directory, if there is one, and increments the `mnt_count`. Finally the value in `mount_lock` is checked against the old value. If there is no change, then the crossing was safe. If there was a change, the `mnt_count` is decremented and the whole process is retried.

When walking up the tree (towards the root) by following a ”..” link, a little more care is needed. In this case the seqlock (which contains both a counter and a spinlock) is fully locked to prevent any changes to any mount points while stepping up. This locking is needed to stabilize the link to the mounted-on dentry, which the refcount on the mount itself doesn't ensure.

`mount_lock` is also used to detect and defend against potential attacks against `LOOKUP_BENEATH` and `LOOKUP_IN_ROOT` when resolving ”..” (where the parent directory is moved outside the root, bypassing the `path_equal()` check). If `mount_lock` is updated during the lookup and the path encounters a ”..”, a potential attack occurred and `handle_dots()` will bail out with `-EAGAIN`.

### RCU

Finally the global (but extremely lightweight) RCU read lock is held from time to time to ensure certain data structures don't get freed unexpectedly.

In particular it is held while scanning chains in the dcache hash table, and the mount point hash table.

### Bringing it together with `struct nameidata`

Throughout the process of walking a path, the current status is stored in a `struct nameidata`, ”namei” being the traditional name - dating all the way back to First Edition Unix - of the function that converts a ”name” to an ”inode”. `struct nameidata` contains (among other fields):

#### `struct path path`

A `path` contains a `struct vfsmount` (which is embedded in a `struct mount`) and a `struct dentry`. Together these record the current status of the walk. They start out referring to the starting point (the current working directory, the root directory, or some other directory identified by a file descriptor), and are updated on each step. A reference through `d_lockref` and `mnt_count` is always held.

#### struct qstr last

This is a string together with a length (i.e. *not* `nul` terminated) that is the "next" component in the pathname.

#### int last_type

This is one of `LAST_NORM`, `LAST_ROOT`, `LAST_DOT` or `LAST_DOTDOT`. The `last` field is only valid if the type is `LAST_NORM`.

#### struct path root

This is used to hold a reference to the effective root of the filesystem. Often that reference won't be needed, so this field is only assigned the first time it is used, or when a non-standard root is requested. Keeping a reference in the `nameidata` ensures that only one root is in effect for the entire path walk, even if it races with a `chroot()` system call.

It should be noted that in the case of `LOOKUP_IN_ROOT` or `LOOKUP_BENEATH`, the effective root becomes the directory file descriptor passed to `openat2()` (which exposes these `LOOKUP_` flags).

The root is needed when either of two conditions holds: (1) either the pathname or a symbolic link starts with a "'/'", or (2) a "`..`" component is being handled, since "`..`" from the root must always stay at the root. The value used is usually the current root directory of the calling process. An alternate root can be provided as when `sysctl()` calls `file_open_root()`, and when NFSv4 or Btrfs call `mount_subtree()`. In each case a pathname is being looked up in a very specific part of the filesystem, and the lookup must not be allowed to escape that subtree. It works a bit like a local `chroot()`.

Ignoring the handling of symbolic links, we can now describe the "`link_path_walk()`" function, which handles the lookup of everything except the final component as:

> Given a path (`name`) and a nameidata structure (`nd`), check that the current directory has execute permission and then advance `name` over one component while updating `last_type` and `last`. If that was the final component, then return, otherwise call `walk_component()` and repeat from the top.

`walk_component()` is even easier. If the component is `LAST_DOTS`, it calls `handle_dots()` which does the necessary locking as already described. If it finds a `LAST_NORM` component it first calls "`lookup_fast()`" which only looks in the dcache, but will ask the filesystem to revalidate the result if it is that sort of filesystem. If that doesn't get a good result, it calls "`lookup_slow()`" which takes `i_rwsem`, rechecks the cache, and then asks the filesystem to find a definitive answer.

As the last step of walk_component(), step_into() will be called either directly from walk_component() or from handle_dots(). It calls handle_mounts(), to check and handle mount points, in which a new `struct path` is created containing a counted reference to the new dentry and a reference to the new `vfsmount` which is only counted if it is different from the previous `vfsmount`. Then if there is a symbolic link, step_into() calls pick_link() to deal with it, otherwise it installs the new `struct path` in the `struct nameidata`, and drops the unneeded references.

This "hand-over-hand" sequencing of getting a reference to the new dentry before dropping the reference to the previous dentry may seem obvious, but is worth pointing out so that we will recognize its analogue in the "RCU-walk" version.

### Handling the final component

`link_path_walk()` only walks as far as setting `nd->last` and `nd->last_type` to refer to the final component of the path. It does not call `walk_component()` that last time. Handling that final component remains for the caller to sort out. Those callers are path_lookupat(), path_parentat() and path_openat() each of which handles the differing requirements of different system calls.

`path_parentat()` is clearly the simplest - it just wraps a little bit of housekeeping around `link_path_walk()` and returns the parent directory and final component to the caller. The caller will be either aiming to create a name (via `filename_create()`) or remove or rename a name (in which case `user_path_parent()` is used). They will use `i_rwsem` to exclude other changes while they validate and then perform their operation.

`path_lookupat()` is nearly as simple - it is used when an existing object is wanted such as by `stat()` or `chmod()`. It essentially just calls `walk_component()` on the final component through a call to `lookup_last()`. `path_lookupat()` returns just the final dentry. It is worth noting that when flag `LOOKUP_MOUNTPOINT` is set, path_lookupat() will unset LOOKUP_JUMPED in nameidata so that in the subsequent path traversal d_weak_revalidate() won't be called. This is important when unmounting a filesystem that is inaccessible, such as one provided by a dead NFS server.

Finally `path_openat()` is used for the `open()` system call; it contains, in support functions starting with "open_last_lookups()", all the complexity needed to handle the different subtleties of O_CREAT (with or without O_EXCL), final "`/`" characters, and trailing symbolic links. We will revisit this in the final part of this series, which focuses on those symbolic links. "open_last_lookups()" will sometimes, but not always, take `i_rwsem`, depending on what it finds.

Each of these, or the functions which call them, need to be alert to the possibility that the final component is not `LAST_NORM`. If the goal of the lookup is to create something, then any value for `last_type` other than `LAST_NORM` will result in an error. For example if `path_parentat()` reports `LAST_DOTDOT`, then the caller won't try to create that name. They also check for trailing slashes by testing `last.name[last.len]`. If there is any character beyond the final component, it must be a trailing slash.

### Revalidation and automounts

Apart from symbolic links, there are only two parts of the "REF-walk" process not yet covered. One is the handling of stale cache entries and the other is automounts.

On filesystems that require it, the lookup routines will call the `->d_revalidate()` dentry method to ensure that the cached information is current. This will often confirm validity or update a few details from a server. In some cases it may find that there has been change further up the path and that something that was thought to be valid previously isn't really. When this happens the lookup of the whole path is aborted and retried with the "`LOOKUP_REVAL`" flag set. This forces revalidation to be more thorough. We will see more details of this retry process in the next article.

Automount points are locations in the filesystem where an attempt to lookup a name can trigger changes to how that lookup should be handled, in particular by mounting a filesystem there. These are covered in greater detail in *autofs - how it works* in the Linux documentation tree, but a few notes specifically related to path lookup are in order here.

The Linux VFS has a concept of "managed" dentries. There are three potentially interesting things about these dentries corresponding to three different flags that might be set in

dentry->d_flags:

#### DCACHE_MANAGE_TRANSIT

If this flag has been set, then the filesystem has requested that the d_manage() dentry operation be called before handling any possible mount point. This can perform two particular services:

It can block to avoid races. If an automount point is being unmounted, the d_manage() function will usually wait for that process to complete before letting the new lookup proceed and possibly trigger a new automount.

It can selectively allow only some processes to transit through a mount point. When a server process is managing automounts, it may need to access a directory without triggering normal automount processing. That server process can identify itself to the autofs filesystem, which will then give it a special pass through d_manage() by returning -EISDIR.

#### DCACHE_MOUNTED

This flag is set on every dentry that is mounted on. As Linux supports multiple filesystem namespaces, it is possible that the dentry may not be mounted on in *this* namespace, just in some other. So this flag is seen as a hint, not a promise.

If this flag is set, and d_manage() didn't return -EISDIR, lookup_mnt() is called to examine the mount hash table (honoring the mount_lock described earlier) and possibly return a new vfsmount and a new dentry (both with counted references).

#### DCACHE_NEED_AUTOMOUNT

If d_manage() allowed us to get this far, and lookup_mnt() didn't find a mount point, then this flag causes the d_automount() dentry operation to be called.

The d_automount() operation can be arbitrarily complex and may communicate with server processes etc. but it should ultimately either report that there was an error, that there was nothing to mount, or should provide an updated struct path with new dentry and vfsmount.

In the latter case, finish_automount() will be called to safely install the new mount point into the mount table.

There is no new locking of import here and it is important that no locks (only counted references) are held over this processing due to the very real possibility of extended delays. This will become more important next time when we examine RCU-walk which is particularly sensitive to delays.

### 1.2.2 RCU-walk - faster pathname lookup in Linux

RCU-walk is another algorithm for performing pathname lookup in Linux. It is in many ways similar to REF-walk and the two share quite a bit of code. The significant difference in RCU-walk is how it allows for the possibility of concurrent access.

We noted that REF-walk is complex because there are numerous details and special cases. RCU-walk reduces this complexity by simply refusing to handle a number of cases -- it instead falls back to REF-walk. The difficulty with RCU-walk comes from a different direction: unfamiliarity. The locking rules when depending on RCU are quite different from traditional locking, so we will spend a little extra time when we come to those.

#### Clear demarcation of roles

The easiest way to manage concurrency is to forcibly stop any other thread from changing the data structures that a given thread is looking at. In cases where no other thread would even think of changing the data and lots of different threads want to read at the same time, this can be very costly. Even when using locks that permit multiple concurrent readers, the simple act of updating the count of the number of current readers can impose an unwanted cost. So the goal when reading a shared data structure that no other process is changing is to avoid writing anything to memory at all. Take no locks, increment no counts, leave no footprints.

The REF-walk mechanism already described certainly doesn't follow this principle, but then it is really designed to work when there may well be other threads modifying the data. RCU-walk, in contrast, is designed for the common situation where there are lots of frequent readers and only occasional writers. This may not be common in all parts of the filesystem tree, but in many parts it will be. For the other parts it is important that RCU-walk can quickly fall back to using REF-walk.

Pathname lookup always starts in RCU-walk mode but only remains there as long as what it is looking for is in the cache and is stable. It dances lightly down the cached filesystem image, leaving no footprints and carefully watching where it is, to be sure it doesn't trip. If it notices that something has changed or is changing, or if something isn't in the cache, then it tries to stop gracefully and switch to REF-walk.

This stopping requires getting a counted reference on the current `vfsmount` and `dentry`, and ensuring that these are still valid - that a path walk with REF-walk would have found the same entries. This is an invariant that RCU-walk must guarantee. It can only make decisions, such as selecting the next step, that are decisions which REF-walk could also have made if it were walking down the tree at the same time. If the graceful stop succeeds, the rest of the path is processed with the reliable, if slightly sluggish, REF-walk. If RCU-walk finds it cannot stop gracefully, it simply gives up and restarts from the top with REF-walk.

This pattern of ”try RCU-walk, if that fails try REF-walk” can be clearly seen in functions like filename_lookup(), filename_parentat(), do_filp_open(), and do_file_open_root(). These four correspond roughly to the three `path_*()` functions we met earlier, each of which calls `link_path_walk()`. The `path_*()` functions are called using different mode flags until a mode is found which works. They are first called with `LOOKUP_RCU` set to request ”RCU-walk”. If that fails with the error `ECHILD` they are called again with no special flag to request ”REF-walk”. If either of those report the error `ESTALE` a final attempt is made with `LOOKUP_REVAL` set (and no `LOOKUP_RCU`) to ensure that entries found in the cache are forcibly revalidated - normally entries are only revalidated if the filesystem determines that they are too old to trust.

The LOOKUP_RCU attempt may drop that flag internally and switch to REF-walk, but will never then try to switch back to RCU-walk. Places that trip up RCU-walk are much more likely to be near the leaves and so it is very unlikely that there will be much, if any, benefit from switching back.

### RCU and seqlocks: fast and light

RCU is, unsurprisingly, critical to RCU-walk mode. The rcu_read_lock() is held for the entire time that RCU-walk is walking down a path. The particular guarantee it provides is that the key data structures - dentries, inodes, super_blocks, and mounts - will not be freed while the lock is held. They might be unlinked or invalidated in one way or another, but the memory will not be repurposed so values in various fields will still be meaningful. This is the only guarantee that RCU provides; everything else is done using seqlocks.

As we saw above, REF-walk holds a counted reference to the current dentry and the current vfsmount, and does not release those references before taking references to the ”next” dentry or vfsmount. It also sometimes takes the d_lock spinlock. These references and locks are taken to prevent certain changes from happening. RCU-walk must not take those references or locks and so cannot prevent such changes. Instead, it checks to see if a change has been made, and aborts or retries if it has.

To preserve the invariant mentioned above (that RCU-walk may only make decisions that REF-walk could have made), it must make the checks at or near the same places that REF-walk holds the references. So, when REF-walk increments a reference count or takes a spinlock, RCU-walk samples the status of a seqlock using read_seqcount_begin() or a similar function. When REF-walk decrements the count or drops the lock, RCU-walk checks if the sampled status is still valid using read_seqcount_retry() or similar.

However, there is a little bit more to seqlocks than that. If RCU-walk accesses two different fields in a seqlock-protected structure, or accesses the same field twice, there is no a priori guarantee of any consistency between those accesses. When consistency is needed - which it usually is - RCU-walk must take a copy and then use read_seqcount_retry() to validate that copy.

read_seqcount_retry() not only checks the sequence number, but also imposes a memory barrier so that no memory-read instruction from *before* the call can be delayed until *after* the call, either by the CPU or by the compiler. A simple example of this can be seen in slow_dentry_cmp() which, for filesystems which do not use simple byte-wise name equality, calls into the filesystem to compare a name against a dentry. The length and name pointer are copied into local variables, then read_seqcount_retry() is called to confirm the two are consistent, and only then is ->d_compare() called. When standard filename comparison is used, dentry_cmp() is called instead. Notably it does *not* use read_seqcount_retry(), but instead has a large comment explaining why the consistency guarantee isn't necessary. A subsequent read_seqcount_retry() will be sufficient to catch any problem that could occur at this point.

With that little refresher on seqlocks out of the way we can look at the bigger picture of how RCU-walk uses seqlocks.

### mount_lock and nd->m_seq

We already met the mount_lock seqlock when REF-walk used it to ensure that crossing a mount point is performed safely. RCU-walk uses it for that too, but for quite a bit more.

Instead of taking a counted reference to each vfsmount as it descends the tree, RCU-walk samples the state of mount_lock at the start of the walk and stores this initial sequence number in the struct nameidata in the m_seq field. This one lock and one sequence number are used to validate all accesses to all vfsmounts, and all mount point crossings. As changes to the mount table are relatively rare, it is reasonable to fall back on REF-walk any time that any ”mount” or ”unmount” happens.

m_seq is checked (using read_seqretry()) at the end of an RCU-walk sequence, whether switching to REF-walk for the rest of the path or when the end of the path is reached. It is also checked when stepping down over a mount point (in __follow_mount_rcu()) or up (in follow_dotdot_rcu()). If it is ever found to have changed, the whole RCU-walk sequence is aborted and the path is processed again by REF-walk.

If RCU-walk finds that mount_lock hasn't changed then it can be sure that, had REF-walk taken counted references on each vfsmount, the results would have been the same. This ensures the invariant holds, at least for vfsmount structures.

### dentry->d_seq and nd->seq

In place of taking a count or lock on d_reflock, RCU-walk samples the per-dentry d_seq seqlock, and stores the sequence number in the seq field of the nameidata structure, so nd->seq should always be the current sequence number of nd->dentry. This number needs to be revalidated after copying, and before using, the name, parent, or inode of the dentry.

The handling of the name we have already looked at, and the parent is only accessed in follow_dotdot_rcu() which fairly trivially follows the required pattern, though it does so for three different cases.

When not at a mount point, d_parent is followed and its d_seq is collected. When we are at a mount point, we instead follow the mnt->mnt_mountpoint link to get a new dentry and collect its d_seq. Then, after finally finding a d_parent to follow, we must check if we have landed on a mount point and, if so, must find that mount point and follow the mnt->mnt_root link. This would imply a somewhat unusual, but certainly possible, circumstance where the starting point of the path lookup was in part of the filesystem that was mounted on, and so not visible from the root.

The inode pointer, stored in ->d_inode, is a little more interesting. The inode will always need to be accessed at least twice, once to determine if it is NULL and once to verify access permissions. Symlink handling requires a validated inode pointer too. Rather than revalidating on each access, a copy is made on the first access and it is stored in the inode field of nameidata from where it can be safely accessed without further validation.

lookup_fast() is the only lookup routine that is used in RCU-mode, lookup_slow() being too slow and requiring locks. It is in lookup_fast() that we find the important ”hand over hand” tracking of the current dentry.

The current dentry and current seq number are passed to __d_lookup_rcu() which, on success, returns a new dentry and a new seq number. lookup_fast() then copies the inode pointer and revalidates the new seq number. It then validates the old dentry with the old seq

number one last time and only then continues. This process of getting the `seq` number of the new dentry and then checking the `seq` number of the old exactly mirrors the process of getting a counted reference to the new dentry before dropping that for the old dentry which we saw in REF-walk.

### No `inode->i_rwsem` or even `rename_lock`

A semaphore is a fairly heavyweight lock that can only be taken when it is permissible to sleep. As `rcu_read_lock()` forbids sleeping, `inode->i_rwsem` plays no role in RCU-walk. If some other thread does take `i_rwsem` and modifies the directory in a way that RCU-walk needs to notice, the result will be either that RCU-walk fails to find the dentry that it is looking for, or it will find a dentry which `read_seqretry()` won't validate. In either case it will drop down to REF-walk mode which can take whatever locks are needed.

Though `rename_lock` could be used by RCU-walk as it doesn't require any sleeping, RCU-walk doesn't bother. REF-walk uses `rename_lock` to protect against the possibility of hash chains in the dcache changing while they are being searched. This can result in failing to find something that actually is there. When RCU-walk fails to find something in the dentry cache, whether it is really there or not, it already drops down to REF-walk and tries again with appropriate locking. This neatly handles all cases, so adding extra checks on rename_lock would bring no significant value.

### `unlazy_walk()` and `complete_walk()`

That "dropping down to REF-walk" typically involves a call to `unlazy_walk()`, so named because "RCU-walk" is also sometimes referred to as "lazy walk". `unlazy_walk()` is called when following the path down to the current vfsmount/dentry pair seems to have proceeded successfully, but the next step is problematic. This can happen if the next name cannot be found in the dcache, if permission checking or name revalidation couldn't be achieved while the `rcu_read_lock()` is held (which forbids sleeping), if an automount point is found, or in a couple of cases involving symlinks. It is also called from `complete_walk()` when the lookup has reached the final component, or the very end of the path, depending on which particular flavor of lookup is used.

Other reasons for dropping out of RCU-walk that do not trigger a call to `unlazy_walk()` are when some inconsistency is found that cannot be handled immediately, such as `mount_lock` or one of the `d_seq` seqlocks reporting a change. In these cases the relevant function will return `-ECHILD` which will percolate up until it triggers a new attempt from the top using REF-walk.

For those cases where `unlazy_walk()` is an option, it essentially takes a reference on each of the pointers that it holds (vfsmount, dentry, and possibly some symbolic links) and then verifies that the relevant seqlocks have not been changed. If there have been changes, it, too, aborts with `-ECHILD`, otherwise the transition to REF-walk has been a success and the lookup process continues.

Taking a reference on those pointers is not quite as simple as just incrementing a counter. That works to take a second reference if you already have one (often indirectly through another object), but it isn't sufficient if you don't actually have a counted reference at all. For `dentry->d_lockref`, it is safe to increment the reference counter to get a reference unless it has been explicitly marked as "dead" which involves setting the counter to `-128`. `lockref_get_not_dead()` achieves this.

For `mnt->mnt_count` it is safe to take a reference as long as `mount_lock` is then used to validate the reference. If that validation fails, it may *not* be safe to just drop that reference in the standard way of calling `mnt_put()` - an unmount may have progressed too far. So the code in `legitimize_mnt()`, when it finds that the reference it got might not be safe, checks the `MNT_SYNC_UMOUNT` flag to determine if a simple `mnt_put()` is correct, or if it should just decrement the count and pretend none of this ever happened.

### Taking care in filesystems

RCU-walk depends almost entirely on cached information and often will not call into the filesystem at all. However there are two places, besides the already-mentioned component-name comparison, where the file system might be included in RCU-walk, and it must know to be careful.

If the filesystem has non-standard permission-checking requirements - such as a networked filesystem which may need to check with the server - the `i_op->permission` interface might be called during RCU-walk. In this case an extra "`MAY_NOT_BLOCK`" flag is passed so that it knows not to sleep, but to return `-ECHILD` if it cannot complete promptly. `i_op->permission` is given the inode pointer, not the dentry, so it doesn't need to worry about further consistency checks. However if it accesses any other filesystem data structures, it must ensure they are safe to be accessed with only the `rcu_read_lock()` held. This typically means they must be freed using `kfree_rcu()` or similar.

If the filesystem may need to revalidate dcache entries, then `d_op->d_revalidate` may be called in RCU-walk too. This interface *is* passed the dentry but does not have access to the `inode` or the `seq` number from the `nameidata`, so it needs to be extra careful when accessing fields in the dentry. This "extra care" typically involves using READ_ONCE() to access fields, and verifying the result is not NULL before using it. This pattern can be seen in `nfs_lookup_revalidate()`.

### A pair of patterns

In various places in the details of REF-walk and RCU-walk, and also in the big picture, there are a couple of related patterns that are worth being aware of.

The first is "try quickly and check, if that fails try slowly". We can see that in the high-level approach of first trying RCU-walk and then trying REF-walk, and in places where `unlazy_walk()` is used to switch to REF-walk for the rest of the path. We also saw it earlier in `dget_parent()` when following a "`..`" link. It tries a quick way to get a reference, then falls back to taking locks if needed.

The second pattern is "try quickly and check, if that fails try again - repeatedly". This is seen with the use of `rename_lock` and `mount_lock` in REF-walk. RCU-walk doesn't make use of this pattern - if anything goes wrong it is much safer to just abort and try a more sedate approach.

The emphasis here is "try quickly and check". It should probably be "try quickly *and carefully*, then check". The fact that checking is needed is a reminder that the system is dynamic and only a limited number of things are safe at all. The most likely cause of errors in this whole process is assuming something is safe when in reality it isn't. Careful consideration of what exactly guarantees the safety of each access is sometimes necessary.

### 1.2.3 A walk among the symlinks

There are several basic issues that we will examine to understand the handling of symbolic links: the symlink stack, together with cache lifetimes, will help us understand the overall recursive handling of symlinks and lead to the special care needed for the final component. Then a consideration of access-time updates and summary of the various flags controlling lookup will finish the story.

#### The symlink stack

There are only two sorts of filesystem objects that can usefully appear in a path prior to the final component: directories and symlinks. Handling directories is quite straightforward: the new directory simply becomes the starting point at which to interpret the next component on the path. Handling symbolic links requires a bit more work.

Conceptually, symbolic links could be handled by editing the path. If a component name refers to a symbolic link, then that component is replaced by the body of the link and, if that body starts with a '/', then all preceding parts of the path are discarded. This is what the "`readlink -f`" command does, though it also edits out "`.`" and "`..`" components.

Directly editing the path string is not really necessary when looking up a path, and discarding early components is pointless as they aren't looked at anyway. Keeping track of all remaining components is important, but they can of course be kept separately; there is no need to concatenate them. As one symlink may easily refer to another, which in turn can refer to a third, we may need to keep the remaining components of several paths, each to be processed when the preceding ones are completed. These path remnants are kept on a stack of limited size.

There are two reasons for placing limits on how many symlinks can occur in a single path lookup. The most obvious is to avoid loops. If a symlink referred to itself either directly or through intermediaries, then following the symlink can never complete successfully - the error `ELOOP` must be returned. Loops can be detected without imposing limits, but limits are the simplest solution and, given the second reason for restriction, quite sufficient.

The second reason was outlined recently by Linus:

> Because it's a latency and DoS issue too. We need to react well to true loops, but also to "very deep" non-loops. It's not about memory use, it's about users triggering unreasonable CPU resources.

Linux imposes a limit on the length of any pathname: `PATH_MAX`, which is 4096. There are a number of reasons for this limit; not letting the kernel spend too much time on just one path is one of them. With symbolic links you can effectively generate much longer paths so some sort of limit is needed for the same reason. Linux imposes a limit of at most 40 (MAXSYMLINKS) symlinks in any one path lookup. It previously imposed a further limit of eight on the maximum depth of recursion, but that was raised to 40 when a separate stack was implemented, so there is now just the one limit.

The `nameidata` structure that we met in an earlier article contains a small stack that can be used to store the remaining part of up to two symlinks. In many cases this will be sufficient. If it isn't, a separate stack is allocated with room for 40 symlinks. Pathname lookup will never exceed that stack as, once the 40th symlink is detected, an error is returned.

It might seem that the name remnants are all that needs to be stored on this stack, but we need a bit more. To see that, we need to move on to cache lifetimes.

### Storage and lifetime of cached symlinks

Like other filesystem resources, such as inodes and directory entries, symlinks are cached by Linux to avoid repeated costly access to external storage. It is particularly important for RCU-walk to be able to find and temporarily hold onto these cached entries, so that it doesn't need to drop down into REF-walk.

While each filesystem is free to make its own choice, symlinks are typically stored in one of two places. Short symlinks are often stored directly in the inode. When a filesystem allocates a `struct inode` it typically allocates extra space to store private data (a common object-oriented design pattern in the kernel). This will sometimes include space for a symlink. The other common location is in the page cache, which normally stores the content of files. The pathname in a symlink can be seen as the content of that symlink and can easily be stored in the page cache just like file content.

When neither of these is suitable, the next most likely scenario is that the filesystem will allocate some temporary memory and copy or construct the symlink content into that memory whenever it is needed.

When the symlink is stored in the inode, it has the same lifetime as the inode which, itself, is protected by RCU or by a counted reference on the dentry. This means that the mechanisms that pathname lookup uses to access the dcache and icache (inode cache) safely are quite sufficient for accessing some cached symlinks safely. In these cases, the `i_link` pointer in the inode is set to point to wherever the symlink is stored and it can be accessed directly whenever needed.

When the symlink is stored in the page cache or elsewhere, the situation is not so straightforward. A reference on a dentry or even on an inode does not imply any reference on cached pages of that inode, and even an `rcu_read_lock()` is not sufficient to ensure that a page will not disappear. So for these symlinks the pathname lookup code needs to ask the filesystem to provide a stable reference and, significantly, needs to release that reference when it is finished with it.

Taking a reference to a cache page is often possible even in RCU-walk mode. It does require making changes to memory, which is best avoided, but that isn't necessarily a big cost and it is better than dropping out of RCU-walk mode completely. Even filesystems that allocate space to copy the symlink into can use `GFP_ATOMIC` to often successfully allocate memory without the need to drop out of RCU-walk. If a filesystem cannot successfully get a reference in RCU-walk mode, it must return `-ECHILD` and `unlazy_walk()` will be called to return to REF-walk mode in which the filesystem is allowed to sleep.

The place for all this to happen is the `i_op->get_link()` inode method. This is called both in RCU-walk and REF-walk. In RCU-walk the `dentry*` argument is NULL, `->get_link()` can return -ECHILD to drop out of RCU-walk. Much like the `i_op->permission()` method we looked at previously, `->get_link()` would need to be careful that all the data structures it references are safe to be accessed while holding no counted reference, only the RCU lock. A callback `struct delayed_called` will be passed to `->get_link()`: file systems can set their own put_link function and argument through set_delayed_call(). Later on, when VFS wants to put link, it will call do_delayed_call() to invoke that callback function with the argument.

In order for the reference to each symlink to be dropped when the walk completes, whether in RCU-walk or REF-walk, the symlink stack needs to contain, along with the path remnants:

- the `struct path` to provide a reference to the previous path
- the `const char *` to provide a reference to the to previous name

- the `seq` to allow the path to be safely switched from RCU-walk to REF-walk
- the `struct delayed_call` for later invocation.

This means that each entry in the symlink stack needs to hold five pointers and an integer instead of just one pointer (the path remnant). On a 64-bit system, this is about 40 bytes per entry; with 40 entries it adds up to 1600 bytes total, which is less than half a page. So it might seem like a lot, but is by no means excessive.

Note that, in a given stack frame, the path remnant (`name`) is not part of the symlink that the other fields refer to. It is the remnant to be followed once that symlink has been fully parsed.

### Following the symlink

The main loop in `link_path_walk()` iterates seamlessly over all components in the path and all of the non-final symlinks. As symlinks are processed, the `name` pointer is adjusted to point to a new symlink, or is restored from the stack, so that much of the loop doesn't need to notice. Getting this `name` variable on and off the stack is very straightforward; pushing and popping the references is a little more complex.

When a symlink is found, walk_component() calls pick_link() via step_into() which returns the link from the filesystem. Providing that operation is successful, the old path `name` is placed on the stack, and the new value is used as the `name` for a while. When the end of the path is found (i.e. `*name` is `'\0'`) the old `name` is restored off the stack and path walking continues.

Pushing and popping the reference pointers (inode, cookie, etc.) is more complex in part because of the desire to handle tail recursion. When the last component of a symlink itself points to a symlink, we want to pop the symlink-just-completed off the stack before pushing the symlink-just-found to avoid leaving empty path remnants that would just get in the way.

It is most convenient to push the new symlink references onto the stack in `walk_component()` immediately when the symlink is found; `walk_component()` is also the last piece of code that needs to look at the old symlink as it walks that last component. So it is quite convenient for `walk_component()` to release the old symlink and pop the references just before pushing the reference information for the new symlink. It is guided in this by three flags: `WALK_NOFOLLOW` which forbids it from following a symlink if it finds one, `WALK_MORE` which indicates that it is yet too early to release the current symlink, and `WALK_TRAILING` which indicates that it is on the final component of the lookup, so we will check userspace flag `LOOKUP_FOLLOW` to decide whether follow it when it is a symlink and call `may_follow_link()` to check if we have privilege to follow it.

### Symlinks with no final component

A pair of special-case symlinks deserve a little further explanation. Both result in a new `struct path` (with mount and dentry) being set up in the `nameidata`, and result in pick_link() returning `NULL`.

The more obvious case is a symlink to "`/`". All symlinks starting with "`/`" are detected in pick_link() which resets the `nameidata` to point to the effective filesystem root. If the symlink only contains "`/`" then there is nothing more to do, no components at all, so `NULL` is returned to indicate that the symlink can be released and the stack frame discarded.

The other case involves things in `/proc` that look like symlinks but aren't really (and are therefore commonly referred to as "magic-links"):

```
$ ls -l /proc/self/fd/1
lrwx------ 1 neilb neilb 64 Jun 13 10:19 /proc/self/fd/1 -> /dev/pts/4
```

Every open file descriptor in any process is represented in `/proc` by something that looks like a symlink. It is really a reference to the target file, not just the name of it. When you `readlink` these objects you get a name that might refer to the same file - unless it has been unlinked or mounted over. When `walk_component()` follows one of these, the `->get_link()` method in "procfs" doesn't return a string name, but instead calls nd_jump_link() which updates the `nameidata` in place to point to that target. `->get_link()` then returns `NULL`. Again there is no final component and pick_link() returns `NULL`.

### Following the symlink in the final component

All this leads to `link_path_walk()` walking down every component, and following all symbolic links it finds, until it reaches the final component. This is just returned in the `last` field of `nameidata`. For some callers, this is all they need; they want to create that `last` name if it doesn't exist or give an error if it does. Other callers will want to follow a symlink if one is found, and possibly apply special handling to the last component of that symlink, rather than just the last component of the original file name. These callers potentially need to call `link_path_walk()` again and again on successive symlinks until one is found that doesn't point to another symlink.

This case is handled by relevant callers of link_path_walk(), such as path_lookupat(), path_openat() using a loop that calls link_path_walk(), and then handles the final component by calling open_last_lookups() or lookup_last(). If it is a symlink that needs to be followed, open_last_lookups() or lookup_last() will set things up properly and return the path so that the loop repeats, calling link_path_walk() again. This could loop as many as 40 times if the last component of each symlink is another symlink.

Of the various functions that examine the final component, open_last_lookups() is the most interesting as it works in tandem with do_open() for opening a file. Part of open_last_lookups() runs with `i_rwsem` held and this part is in a separate function: lookup_open().

Explaining open_last_lookups() and do_open() completely is beyond the scope of this article, but a few highlights should help those interested in exploring the code.

1. Rather than just finding the target file, do_open() is used after open_last_lookup() to open it. If the file was found in the dcache, then `vfs_open()` is used for this. If not, then `lookup_open()` will either call `atomic_open()` (if the filesystem provides it) to combine the final lookup with the open, or will perform the separate `i_op->lookup()` and `i_op->create()` steps directly. In the later case the actual "open" of this newly found or created file will be performed by vfs_open(), just as if the name were found in the dcache.

2. vfs_open() can fail with `-EOPENSTALE` if the cached information wasn't quite current enough. If it's in RCU-walk `-ECHILD` will be returned otherwise `-ESTALE` is returned. When `-ESTALE` is returned, the caller may retry with `LOOKUP_REVAL` flag set.

3. An open with O_CREAT **does** follow a symlink in the final component, unlike other creation system calls (like `mkdir`). So the sequence:

   ```
   ln -s bar /tmp/foo
   echo hello > /tmp/foo
   ```

will create a file called `/tmp/bar`. This is not permitted if `O_EXCL` is set but otherwise is handled for an O_CREAT open much like for a non-creating open: lookup_last() or open_last_lookup() returns a non `NULL` value, and link_path_walk() gets called and the open process continues on the symlink that was found.

### Updating the access time

We previously said of RCU-walk that it would ”take no locks, increment no counts, leave no footprints.” We have since seen that some ”footprints” can be needed when handling symlinks as a counted reference (or even a memory allocation) may be needed. But these footprints are best kept to a minimum.

One other place where walking down a symlink can involve leaving footprints in a way that doesn't affect directories is in updating access times. In Unix (and Linux) every filesystem object has a ”last accessed time”, or ”`atime`”. Passing through a directory to access a file within is not considered to be an access for the purposes of `atime`; only listing the contents of a directory can update its `atime`. Symlinks are different it seems. Both reading a symlink (with `readlink()`) and looking up a symlink on the way to some other destination can update the atime on that symlink.

It is not clear why this is the case; POSIX has little to say on the subject. The clearest statement is that, if a particular implementation updates a timestamp in a place not specified by POSIX, this must be documented ”except that any changes caused by pathname resolution need not be documented”. This seems to imply that POSIX doesn't really care about access-time updates during pathname lookup.

An examination of history shows that prior to Linux 1.3.87, the ext2 filesystem, at least, didn't update atime when following a link. Unfortunately we have no record of why that behavior was changed.

In any case, access time must now be updated and that operation can be quite complex. Trying to stay in RCU-walk while doing it is best avoided. Fortunately it is often permitted to skip the `atime` update. Because `atime` updates cause performance problems in various areas, Linux supports the `relatime` mount option, which generally limits the updates of `atime` to once per day on files that aren't being changed (and symlinks never change once created). Even without `relatime`, many filesystems record `atime` with a one-second granularity, so only one update per second is required.

It is easy to test if an `atime` update is needed while in RCU-walk mode and, if it isn't, the update can be skipped and RCU-walk mode continues. Only when an `atime` update is actually required does the path walk drop down to REF-walk. All of this is handled in the `get_link()` function.

### A few flags

A suitable way to wrap up this tour of pathname walking is to list the various flags that can be stored in the `nameidata` to guide the lookup process. Many of these are only meaningful on the final component, others reflect the current state of the pathname lookup, and some apply restrictions to all path components encountered in the path lookup.

And then there is `LOOKUP_EMPTY`, which doesn't fit conceptually with the others. If this is not set, an empty pathname causes an error very early on. If it is set, empty pathnames are not considered to be an error.

### Global state flags

We have already met two global state flags: LOOKUP_RCU and LOOKUP_REVAL. These select between one of three overall approaches to lookup: RCU-walk, REF-walk, and REF-walk with forced revalidation.

LOOKUP_PARENT indicates that the final component hasn't been reached yet. This is primarily used to tell the audit subsystem the full context of a particular access being audited.

ND_ROOT_PRESET indicates that the root field in the nameidata was provided by the caller, so it shouldn't be released when it is no longer needed.

ND_JUMPED means that the current dentry was chosen not because it had the right name but for some other reason. This happens when following "..", following a symlink to /, crossing a mount point or accessing a "/proc/$PID/fd/$FD" symlink (also known as a "magic link"). In this case the filesystem has not been asked to revalidate the name (with d_revalidate()). In such cases the inode may still need to be revalidated, so d_op->d_weak_revalidate() is called if ND_JUMPED is set when the look completes - which may be at the final component or, when creating, unlinking, or renaming, at the penultimate component.

### Resolution-restriction flags

In order to allow userspace to protect itself against certain race conditions and attack scenarios involving changing path components, a series of flags are available which apply restrictions to all path components encountered during path lookup. These flags are exposed through openat2()'s resolve field.

LOOKUP_NO_SYMLINKS blocks all symlink traversals (including magic-links). This is distinctly different from LOOKUP_FOLLOW, because the latter only relates to restricting the following of trailing symlinks.

LOOKUP_NO_MAGICLINKS blocks all magic-link traversals. Filesystems must ensure that they return errors from nd_jump_link(), because that is how LOOKUP_NO_MAGICLINKS and other magic-link restrictions are implemented.

LOOKUP_NO_XDEV blocks all vfsmount traversals (this includes both bind-mounts and ordinary mounts). Note that the vfsmount which contains the lookup is determined by the first mountpoint the path lookup reaches -- absolute paths start with the vfsmount of /, and relative paths start with the dfd's vfsmount. Magic-links are only permitted if the vfsmount of the path is unchanged.

LOOKUP_BENEATH blocks any path components which resolve outside the starting point of the resolution. This is done by blocking nd_jump_root() as well as blocking ".." if it would jump outside the starting point. rename_lock and mount_lock are used to detect attacks against the resolution of "..". Magic-links are also blocked.

LOOKUP_IN_ROOT resolves all path components as though the starting point were the filesystem root. nd_jump_root() brings the resolution back to the starting point, and ".." at the starting point will act as a no-op. As with LOOKUP_BENEATH, rename_lock and mount_lock are used to detect attacks against ".." resolution. Magic-links are also blocked.

#### Final-component flags

Some of these flags are only set when the final component is being considered. Others are only checked for when considering that final component.

LOOKUP_AUTOMOUNT ensures that, if the final component is an automount point, then the mount is triggered. Some operations would trigger it anyway, but operations like `stat()` deliberately don't. `statfs()` needs to trigger the mount but otherwise behaves a lot like `stat()`, so it sets LOOKUP_AUTOMOUNT, as does "`quotactl()`" and the handling of "`mount --bind`".

LOOKUP_FOLLOW has a similar function to LOOKUP_AUTOMOUNT but for symlinks. Some system calls set or clear it implicitly, while others have API flags such as AT_SYMLINK_FOLLOW and UMOUNT_NOFOLLOW to control it. Its effect is similar to WALK_GET that we already met, but it is used in a different way.

LOOKUP_DIRECTORY insists that the final component is a directory. Various callers set this and it is also set when the final component is found to be followed by a slash.

Finally LOOKUP_OPEN, LOOKUP_CREATE, LOOKUP_EXCL, and LOOKUP_RENAME_TARGET are not used directly by the VFS but are made available to the filesystem and particularly the `->d_revalidate()` method. A filesystem can choose not to bother revalidating too hard if it knows that it will be asked to open or create the file soon. These flags were previously useful for `->lookup()` too but with the introduction of `->atomic_open()` they are less relevant there.

#### End of the road

Despite its complexity, all this pathname lookup code appears to be in good shape - various parts are certainly easier to understand now than even a couple of releases ago. But that doesn't mean it is "finished". As already mentioned, RCU-walk currently only follows symlinks that are stored in the inode so, while it handles many ext4 symlinks, it doesn't help with NFS, XFS, or Btrfs. That support is not likely to be long delayed.

## 1.3 Linux Filesystems API summary

This section contains API-level documentation, mostly taken from the source code itself.

### 1.3.1 The Linux VFS

#### The Filesystem types

enum **positive_aop_returns**

aop return codes with specific semantics

**Constants**

**AOP_WRITEPAGE_ACTIVATE**

Informs the caller that page writeback has completed, that the page is still locked, and should be considered active. The VM uses this hint to return the page to the active list -- it won't be a candidate for writeback again in the near future. Other callers must be careful to unlock the page if they get this return. Returned by writepage();

**AOP_TRUNCATED_PAGE**
The AOP method that was handed a locked page has unlocked it and the page might have been truncated. The caller should back up to acquiring a new page and trying again. The aop will be taking reasonable precautions not to livelock. If the caller held a page reference, it should drop it before retrying. Returned by read_folio().

**Description**

address_space_operation functions return these large constants to indicate special semantics to the caller. These are much larger than the bytes in a page to allow for functions that return the number of bytes operated on in a given page.

struct **address_space**
Contents of a cacheable, mappable object.

**Definition**:

```
struct address_space {
    struct inode            *host;
    struct xarray           i_pages;
    struct rw_semaphore     invalidate_lock;
    gfp_t gfp_mask;
    atomic_t i_mmap_writable;
#ifdef CONFIG_READ_ONLY_THP_FOR_FS;
    atomic_t nr_thps;
#endif;
    struct rb_root_cached   i_mmap;
    unsigned long           nrpages;
    pgoff_t writeback_index;
    const struct address_space_operations *a_ops;
    unsigned long           flags;
    struct rw_semaphore     i_mmap_rwsem;
    errseq_t wb_err;
    spinlock_t i_private_lock;
    struct list_head        i_private_list;
    void *                  i_private_data;
};
```

**Members**

**host**
Owner, either the inode or the block_device.

**i_pages**
Cached pages.

**invalidate_lock**
Guards coherency between page cache contents and file offset->disk block mappings in the filesystem during invalidates. It is also used to block modification of page cache contents through memory mappings.

**gfp_mask**
Memory allocation flags to use for allocating pages.

**i_mmap_writable**
Number of VM_SHARED, VM_MAYWRITE mappings.

**nr_thps**
Number of THPs in the pagecache (non-shmem only).

**i_mmap**
Tree of private and shared mappings.

**nrpages**
Number of page entries, protected by the i_pages lock.

**writeback_index**
Writeback starts here.

**a_ops**
Methods.

**flags**
Error bits and flags (AS_*).

**i_mmap_rwsem**
Protects **i_mmap** and **i_mmap_writable**.

**wb_err**
The most recent error which has occurred.

**i_private_lock**
For use by the owner of the address_space.

**i_private_list**
For use by the owner of the address_space.

**i_private_data**
For use by the owner of the address_space.

struct **file_ra_state**
Track a file's readahead state.

**Definition**:

```
struct file_ra_state {
    pgoff_t start;
    unsigned int size;
    unsigned int async_size;
    unsigned int ra_pages;
    unsigned int mmap_miss;
    loff_t prev_pos;
};
```

**Members**

**start**
Where the most recent readahead started.

**size**
Number of pages read in the most recent readahead.

**async_size**
Numer of pages that were/are not needed immediately and so were/are genuinely ”ahead”. Start next readahead when the first of these pages is accessed.

**ra_pages**
Maximum size of a readahead request, copied from the bdi.

**mmap_miss**
How many mmap accesses missed in the page cache.

**prev_pos**
The last byte in the most recent read request.

**Description**

When this structure is passed to ->readahead(), the ”most recent” readahead means the current readahead.

vfsuid_t **i_uid_into_vfsuid**(struct mnt_idmap *idmap, const struct *inode* *inode)
map an inode's i_uid down according to an idmapping

**Parameters**

**struct mnt_idmap *idmap**
idmap of the mount the inode was found from

**const struct inode *inode**
inode to map

**Return**

whe inode's i_uid mapped down according to **idmap**. If the inode's i_uid has no mapping INVALID_VFSUID is returned.

bool **i_uid_needs_update**(struct mnt_idmap *idmap, const struct iattr *attr, const struct *inode* *inode)
check whether inode's i_uid needs to be updated

**Parameters**

**struct mnt_idmap *idmap**
idmap of the mount the inode was found from

**const struct iattr *attr**
the new attributes of **inode**

**const struct inode *inode**
the inode to update

**Description**

Check whether the $inode's i_uid field needs to be updated taking idmapped mounts into account if the filesystem supports it.

**Return**

true if **inode**'s i_uid field needs to be updated, false if not.

void **i_uid_update**(struct mnt_idmap *idmap, const struct iattr *attr, struct *inode* *inode)
update **inode**'s i_uid field

**Parameters**

**struct mnt_idmap *idmap**
idmap of the mount the inode was found from

**const struct iattr *attr**
the new attributes of **inode**

**struct inode *inode**
the inode to update

**Description**

Safely update **inode**'s i_uid field translating the vfsuid of any idmapped mount into the filesystem kuid.

vfsgid_t **i_gid_into_vfsgid**(struct mnt_idmap *idmap, const struct *inode* *inode)
map an inode's i_gid down according to an idmapping

**Parameters**

**struct mnt_idmap *idmap**
idmap of the mount the inode was found from

**const struct inode *inode**
inode to map

**Return**

the inode's i_gid mapped down according to **idmap**. If the inode's i_gid has no mapping INVALID_VFSGID is returned.

bool **i_gid_needs_update**(struct mnt_idmap *idmap, const struct iattr *attr, const struct *inode* *inode)
check whether inode's i_gid needs to be updated

**Parameters**

**struct mnt_idmap *idmap**
idmap of the mount the inode was found from

**const struct iattr *attr**
the new attributes of **inode**

**const struct inode *inode**
the inode to update

**Description**

Check whether the $inode's i_gid field needs to be updated taking idmapped mounts into account if the filesystem supports it.

**Return**

true if **inode**'s i_gid field needs to be updated, false if not.

void **i_gid_update**(struct mnt_idmap *idmap, const struct iattr *attr, struct *inode* *inode)
update **inode**'s i_gid field

**Parameters**

**struct mnt_idmap *idmap**
idmap of the mount the inode was found from

**const struct iattr *attr**
the new attributes of **inode**

**struct inode *inode**
: the inode to update

**Description**

Safely update **inode**'s i_gid field translating the vfsgid of any idmapped mount into the filesystem kgid.

void **inode_fsuid_set**(struct *inode* *inode, struct mnt_idmap *idmap)
: initialize inode's i_uid field with callers fsuid

**Parameters**

**struct inode *inode**
: inode to initialize

**struct mnt_idmap *idmap**
: idmap of the mount the inode was found from

**Description**

Initialize the i_uid field of **inode**. If the inode was found/created via an idmapped mount map the caller's fsuid according to **idmap**.

void **inode_fsgid_set**(struct *inode* *inode, struct mnt_idmap *idmap)
: initialize inode's i_gid field with callers fsgid

**Parameters**

**struct inode *inode**
: inode to initialize

**struct mnt_idmap *idmap**
: idmap of the mount the inode was found from

**Description**

Initialize the i_gid field of **inode**. If the inode was found/created via an idmapped mount map the caller's fsgid according to **idmap**.

bool **fsuidgid_has_mapping**(struct super_block *sb, struct mnt_idmap *idmap)
: check whether caller's fsuid/fsgid is mapped

**Parameters**

**struct super_block *sb**
: the superblock we want a mapping in

**struct mnt_idmap *idmap**
: idmap of the relevant mount

**Description**

Check whether the caller's fsuid and fsgid have a valid mapping in the s_user_ns of the superblock **sb**. If the caller is on an idmapped mount map the caller's fsuid and fsgid according to the **idmap** first.

**Return**

true if fsuid and fsgid is mapped, false if not.

struct timespec64 **inode_set_ctime**(struct *inode* *inode, time64_t sec, long nsec)
set the ctime in the inode

**Parameters**

**struct inode *inode**
inode in which to set the ctime

**time64_t sec**
tv_sec value to set

**long nsec**
tv_nsec value to set

**Description**

Set the ctime in **inode** to { **sec**, **nsec** }

int **__sb_write_started**(const struct super_block *sb, int level)
check if sb freeze level is held

**Parameters**

**const struct super_block *sb**
the super we write to

**int level**
the freeze level

**Description**

- > 0 - sb freeze level is held
- 0 - sb freeze level is not held
- < 0 - !CONFIG_LOCKDEP/LOCK_STATE_UNKNOWN

bool **sb_write_started**(const struct super_block *sb)
check if SB_FREEZE_WRITE is held

**Parameters**

**const struct super_block *sb**
the super we write to

**Description**

May be false positive with !CONFIG_LOCKDEP/LOCK_STATE_UNKNOWN.

bool **sb_write_not_started**(const struct super_block *sb)
check if SB_FREEZE_WRITE is not held

**Parameters**

**const struct super_block *sb**
the super we write to

**Description**

May be false positive with !CONFIG_LOCKDEP/LOCK_STATE_UNKNOWN.

bool **file_write_started**(const struct *file* *file)
check if SB_FREEZE_WRITE is held

**Parameters**

**const struct file *file**
the file we write to

**Description**

May be false positive with !CONFIG_LOCKDEP/LOCK_STATE_UNKNOWN. May be false positive with !S_ISREG, because *file_start_write()* has no effect on !S_ISREG.

bool **file_write_not_started**(const struct *file* *file)
check if SB_FREEZE_WRITE is not held

**Parameters**

**const struct file *file**
the file we write to

**Description**

May be false positive with !CONFIG_LOCKDEP/LOCK_STATE_UNKNOWN. May be false positive with !S_ISREG, because *file_start_write()* has no effect on !S_ISREG.

void **sb_end_write**(struct super_block *sb)
drop write access to a superblock

**Parameters**

**struct super_block *sb**
the super we wrote to

**Description**

Decrement number of writers to the filesystem. Wake up possible waiters wanting to freeze the filesystem.

void **sb_end_pagefault**(struct super_block *sb)
drop write access to a superblock from a page fault

**Parameters**

**struct super_block *sb**
the super we wrote to

**Description**

Decrement number of processes handling write page fault to the filesystem. Wake up possible waiters wanting to freeze the filesystem.

void **sb_end_intwrite**(struct super_block *sb)
drop write access to a superblock for internal fs purposes

**Parameters**

**struct super_block *sb**
the super we wrote to

**Description**

Decrement fs-internal number of writers to the filesystem. Wake up possible waiters wanting to freeze the filesystem.

void **sb_start_write**(struct super_block *sb)

get write access to a superblock

**Parameters**

**struct super_block *sb**
the super we write to

**Description**

When a process wants to write data or metadata to a file system (i.e. dirty a page or an inode), it should embed the operation in a *sb_start_write()* - *sb_end_write()* pair to get exclusion against file system freezing. This function increments number of writers preventing freezing. If the file system is already frozen, the function waits until the file system is thawed.

Since freeze protection behaves as a lock, users have to preserve ordering of freeze protection and other filesystem locks. Generally, freeze protection should be the outermost lock. In particular, we have:

**sb_start_write**
-> i_mutex (write path, truncate, directory ops, ...) -> s_umount (freeze_super, thaw_super)

void **sb_start_pagefault**(struct super_block *sb)

get write access to a superblock from a page fault

**Parameters**

**struct super_block *sb**
the super we write to

**Description**

When a process starts handling write page fault, it should embed the operation into *sb_start_pagefault()* - *sb_end_pagefault()* pair to get exclusion against file system freezing. This is needed since the page fault is going to dirty a page. This function increments number of running page faults preventing freezing. If the file system is already frozen, the function waits until the file system is thawed.

Since page fault freeze protection behaves as a lock, users have to preserve ordering of freeze protection and other filesystem locks. It is advised to put *sb_start_pagefault()* close to mmap_lock in lock ordering. Page fault handling code implies lock dependency:

**mmap_lock**
-> sb_start_pagefault

void **sb_start_intwrite**(struct super_block *sb)

get write access to a superblock for internal fs purposes

**Parameters**

**struct super_block *sb**
the super we write to

### Description

This is the third level of protection against filesystem freezing. It is free for use by a filesystem. The only requirement is that it must rank below sb_start_pagefault.

For example filesystem can call *sb_start_intwrite()* when starting a transaction which somewhat eases handling of freezing for internal sources of filesystem changes (internal fs threads, discarding preallocation on file close, etc.).

struct **renamedata**

contains all information required for renaming

**Definition**:

```
struct renamedata {
    struct mnt_idmap *old_mnt_idmap;
    struct inode *old_dir;
    struct dentry *old_dentry;
    struct mnt_idmap *new_mnt_idmap;
    struct inode *new_dir;
    struct dentry *new_dentry;
    struct inode **delegated_inode;
    unsigned int flags;
};
```

### Members

**old_mnt_idmap**
idmap of the old mount the inode was found from

**old_dir**
parent of source

**old_dentry**
source

**new_mnt_idmap**
idmap of the new mount the inode was found from

**new_dir**
parent of destination

**new_dentry**
destination

**delegated_inode**
returns an inode needing a delegation break

**flags**
rename flags

enum **freeze_holder**

holder of the freeze

### Constants

**FREEZE_HOLDER_KERNEL**
kernel wants to freeze or thaw filesystem

**FREEZE_HOLDER_USERSPACE**
userspace wants to freeze or thaw filesystem

**FREEZE_MAY_NEST**
whether nesting freeze and thaw requests is allowed

**Description**

Indicate who the owner of the freeze or thaw request is and whether the freeze needs to be exclusive or can nest. Without **FREEZE_MAY_NEST**, multiple freeze and thaw requests from the same holder aren't allowed. It is however allowed to hold a single **FREEZE_HOLDER_USERSPACE** and a single **FREEZE_HOLDER_KERNEL** freeze at the same time. This is relied upon by some filesystems during online repair or similar.

bool **is_idmapped_mnt**(const struct vfsmount *mnt)
check whether a mount is mapped

**Parameters**

**const struct vfsmount *mnt**
the mount to check

**Description**

If **mnt** has an non **nop_mnt_idmap** attached to it then **mnt** is mapped.

**Return**

true if mount is mapped, false if not.

void **file_start_write**(struct *file* *file)
get write access to a superblock for regular file io

**Parameters**

**struct file *file**
the file we want to write to

**Description**

This is a variant of *sb_start_write()* which is a noop on non-regualr file. Should be matched with a call to *file_end_write()*.

void **file_end_write**(struct *file* *file)
drop write access to a superblock of a regular file

**Parameters**

**struct file *file**
the file we wrote to

**Description**

Should be matched with a call to *file_start_write()*.

void **kiocb_start_write**(struct kiocb *iocb)
get write access to a superblock for async file io

**Parameters**

**struct kiocb *iocb**
the io context we want to submit the write with

**Description**

This is a variant of *sb_start_write()* for async io submission. Should be matched with a call to *kiocb_end_write()*.

void **kiocb_end_write**(struct kiocb *iocb)

drop write access to a superblock after async file io

**Parameters**

**struct kiocb *iocb**

the io context we sumbitted the write with

**Description**

Should be matched with a call to *kiocb_start_write()*.

void **inode_dio_begin**(struct *inode* *inode)

signal start of a direct I/O requests

**Parameters**

**struct inode *inode**

inode the direct I/O happens on

**Description**

This is called once we've finished processing a direct I/O request, and is used to wake up callers waiting for direct I/O to be quiesced.

void **inode_dio_end**(struct *inode* *inode)

signal finish of a direct I/O requests

**Parameters**

**struct inode *inode**

inode the direct I/O happens on

**Description**

This is called once we've finished processing a direct I/O request, and is used to wake up callers waiting for direct I/O to be quiesced.

## The Directory Cache

void **d_drop**(struct *dentry* *dentry)

drop a dentry

**Parameters**

**struct dentry *dentry**

dentry to drop

**Description**

*d_drop()* unhashes the entry from the parent dentry hashes, so that it won't be found through a VFS lookup any more. Note that this is different from deleting the dentry - d_delete will try to mark the dentry negative if possible, giving a successful _negative_ lookup, while d_drop will just make the cache lookup fail.

*d_drop()* is used mainly for stuff that wants to invalidate a dentry for some reason (NFS timeouts or autofs deletes).

__d_drop requires dentry->d_lock

___d_drop doesn't mark dentry as "unhashed" (dentry->d_hash.pprev will be LIST_POISON2, not NULL).

struct dentry *__d_find_any_alias__(struct *inode* *inode)
find any alias for a given inode

**Parameters**

**struct inode *inode**
inode to find an alias for

**Description**

If any aliases exist for the given inode, take and return a reference for one of them. If no aliases exist, return `NULL`.

struct dentry *__d_find_alias__(struct *inode* *inode)
grab a hashed alias of inode

**Parameters**

**struct inode *inode**
inode in question

**Description**

If inode has a hashed alias, or is a directory and has any alias, acquire the reference to alias and return it. Otherwise return NULL. Notice that if inode is a directory there can be only one alias and it can be unhashed only if it has no children, or if it is the root of a filesystem, or if the directory was renamed and d_revalidate was the first vfs operation to notice.

If the inode has an IS_ROOT, DCACHE_DISCONNECTED alias, then prefer any other hashed alias over that one.

void __shrink_dcache_sb__(struct super_block *sb)
shrink dcache for a superblock

**Parameters**

**struct super_block *sb**
superblock

**Description**

Shrink the dcache for the specified super block. This is used to free the dcache before unmounting a file system.

int __path_has_submounts__(const struct path *parent)
check for mounts over a dentry in the current namespace.

**Parameters**

**const struct path *parent**
path to check.

**Description**

Return true if the parent or its subdirectories contain a mount point in the current namespace.

void **shrink_dcache_parent**(struct dentry *parent)
prune dcache

**Parameters**

**struct dentry *parent**
parent of entries to prune

**Description**

Prune the dcache to remove unused children of the parent dentry.

void **d_invalidate**(struct *dentry* *dentry)
detach submounts, prune dcache, and drop

**Parameters**

**struct dentry *dentry**
dentry to invalidate (aka detach, prune and drop)

struct dentry ***d_alloc**(struct dentry *parent, const struct qstr *name)
allocate a dcache entry

**Parameters**

**struct dentry * parent**
parent of entry to allocate

**const struct qstr *name**
qstr of the name

**Description**

Allocates a dentry. It returns `NULL` if there is insufficient memory available. On a success the dentry is returned. The name passed in is copied and the copy passed in may be reused after this call.

void **d_instantiate**(struct dentry *entry, struct *inode* *inode)
fill in inode information for a dentry

**Parameters**

**struct dentry *entry**
dentry to complete

**struct inode * inode**
inode to attach to this dentry

**Description**

Fill in inode information in the entry.

This turns negative dentries into productive full members of society.

NOTE! This assumes that the inode count has been incremented (or otherwise set) by the caller to indicate that it is now in use by the dcache.

struct dentry *__d_obtain_alias__(struct *inode* *inode)

find or allocate a DISCONNECTED dentry for a given inode

**Parameters**

**struct inode *inode**
inode to allocate the dentry for

**Description**

Obtain a dentry for an inode resulting from NFS filehandle conversion or similar open by handle operations. The returned dentry may be anonymous, or may have a full name (if the inode was already in the cache).

When called on a directory inode, we must ensure that the inode only ever has one dentry. If a dentry is found, that is returned instead of allocating a new one.

On successful return, the reference to the inode has been transferred to the dentry. In case of an error the reference on the inode is released. To make it easier to use in export operations a `NULL` or IS_ERR inode may be passed in and the error will be propagated to the return value, with a `NULL` **inode** replaced by ERR_PTR(-ESTALE).

struct dentry *__d_obtain_root__(struct *inode* *inode)

find or allocate a dentry for a given inode

**Parameters**

**struct inode *inode**
inode to allocate the dentry for

**Description**

Obtain an IS_ROOT dentry for the root of a filesystem.

We must ensure that directory inodes only ever have one dentry. If a dentry is found, that is returned instead of allocating a new one.

On successful return, the reference to the inode has been transferred to the dentry. In case of an error the reference on the inode is released. A `NULL` or IS_ERR inode may be passed in and will be the error will be propagate to the return value, with a `NULL` **inode** replaced by ERR_PTR(-ESTALE).

struct *dentry* *__d_add_ci__(struct *dentry* *dentry, struct *inode* *inode, struct qstr *name)

lookup or allocate new dentry with case-exact name

**Parameters**

**struct dentry *dentry**
the negative dentry that was passed to the parent's lookup func

**struct inode *inode**
the inode case-insensitive lookup has found

**struct qstr *name**
the case-exact name to be associated with the returned dentry

**Description**

This is to avoid filling the dcache with case-insensitive names to the same inode, only the actual correct case is stored in the dcache for case-insensitive filesystems.

For a case-insensitive lookup match and if the case-exact dentry already exists in the dcache, use it and return it.

If no entry exists with the exact case name, allocate new dentry with the exact case, and return the spliced entry.

bool **d_same_name**(const struct *dentry* *dentry, const struct *dentry* *parent, const struct qstr *name)

compare dentry name with case-exact name

**Parameters**

**const struct dentry *dentry**
the negative dentry that was passed to the parent's lookup func

**const struct dentry *parent**
parent dentry

**const struct qstr *name**
the case-exact name to be associated with the returned dentry

**Return**

true if names are same, or false

struct dentry ***d_lookup**(const struct dentry *parent, const struct qstr *name)

search for a dentry

**Parameters**

**const struct dentry *parent**
parent dentry

**const struct qstr *name**
qstr of name we wish to find

**Return**

dentry, or NULL

**Description**

d_lookup searches the children of the parent dentry for the name in question. If the dentry is found its reference count is incremented and the dentry is returned. The caller must use dput to free the entry when it has finished using it. `NULL` is returned if the dentry does not exist.

struct dentry ***d_hash_and_lookup**(struct dentry *dir, struct qstr *name)

hash the qstr then search for a dentry

**Parameters**

**struct dentry *dir**
Directory to search in

**struct qstr *name**
qstr of name we wish to find

**Description**

On lookup failure NULL is returned; on bad name - ERR_PTR(-error)

void **d_delete**(struct *dentry* *dentry)
delete a dentry

**Parameters**

**struct dentry * dentry**
The dentry to delete

**Description**

Turn the dentry into a negative dentry if possible, otherwise remove it from the hash queues so it can be deleted later

void **d_rehash**(struct dentry *entry)
add an entry back to the hash

**Parameters**

**struct dentry * entry**
dentry to add to the hash

**Description**

Adds a dentry to the hash according to its name.

void **d_add**(struct dentry *entry, struct *inode* *inode)
add dentry to hash queues

**Parameters**

**struct dentry *entry**
dentry to add

**struct inode *inode**
The inode to attach to this dentry

**Description**

This adds the entry to the hash queues and initializes **inode**. The entry was actually filled in earlier during *d_alloc()*.

struct dentry ***d_exact_alias**(struct dentry *entry, struct *inode* *inode)
find and hash an exact unhashed alias

**Parameters**

**struct dentry *entry**
dentry to add

**struct inode *inode**
The inode to go with this dentry

**Description**

If an unhashed dentry with the same name/parent and desired inode already exists, hash and return it. Otherwise, return NULL.

Parent directory should be locked.

struct *dentry* ***d_splice_alias**(struct *inode* *inode, struct *dentry* *dentry)
splice a disconnected dentry into the tree if one exists

**Parameters**

**struct inode *inode**
the inode which may have a disconnected dentry

**struct dentry *dentry**
a negative dentry which we want to point to the inode.

**Description**

If inode is a directory and has an IS_ROOT alias, then d_move that in place of the given dentry and return it, else simply d_add the inode to the dentry and return NULL.

If a non-IS_ROOT directory is found, the filesystem is corrupt, and we should error out: directories can't have multiple aliases.

This is needed in the lookup routine of any filesystem that is exportable (via knfsd) so that we can build dcache paths to directories effectively.

If a dentry was found and moved, then it is returned. Otherwise NULL is returned. This matches the expected return value of ->lookup.

Cluster filesystems may call this function with a negative, hashed dentry. In that case, we know that the inode will be a regular file, and also this will only occur during atomic_open. So we need to check for the dentry being already hashed only in the final case.

bool **is_subdir**(struct dentry *new_dentry, struct dentry *old_dentry)
is new dentry a subdirectory of old_dentry

**Parameters**

**struct dentry *new_dentry**
new dentry

**struct dentry *old_dentry**
old dentry

**Description**

Returns true if new_dentry is a subdirectory of the parent (at any depth). Returns false otherwise. Caller must ensure that "new_dentry" is pinned before calling *is_subdir()*

struct *dentry* ***dget_dlock**(struct *dentry* *dentry)
get a reference to a dentry

**Parameters**

**struct dentry *dentry**
dentry to get a reference to

**Description**

Given a live dentry, increment the reference count and return the dentry. Caller must hold **dentry->d_lock**. Making sure that dentry is alive is caller's resonsibility. There are many conditions sufficient to guarantee that; e.g. anything with non-negative refcount is alive, so's anything hashed, anything positive, anyone's parent, etc.

struct *dentry* ***dget**(struct *dentry* *dentry)
get a reference to a dentry

**Parameters**

**struct dentry *dentry**
dentry to get a reference to

**Description**

Given a dentry or NULL pointer increment the reference count if appropriate and return the dentry. A dentry will not be destroyed when it has references. Conversely, a dentry with no references can disappear for any number of reasons, starting with memory pressure. In other words, that primitive is used to clone an existing reference; using it on something with zero refcount is a bug.

**NOTE**

it will spin if **dentry->d_lock** is held. From the deadlock avoidance point of view it is equivalent to spin_lock()/increment refcount/spin_unlock(), so calling it under **dentry->d_lock** is always a bug; so's calling it under ->d_lock on any of its descendents.

int **d_unhashed**(const struct *dentry* *dentry)
is dentry hashed

**Parameters**

**const struct dentry *dentry**
entry to check

**Description**

Returns true if the dentry passed is not currently hashed.

bool **d_really_is_negative**(const struct *dentry* *dentry)
Determine if a dentry is really negative (ignoring fallthroughs)

**Parameters**

**const struct dentry *dentry**
The dentry in question

**Description**

Returns true if the dentry represents either an absent name or a name that doesn't map to an inode (ie. ->d_inode is NULL). The dentry could represent a true miss, a whiteout that isn't represented by a 0,0 chardev or a fallthrough marker in an opaque directory.

Note! (1) This should be used *only* by a filesystem to examine its own dentries. It should not be used to look at some other filesystem's dentries. (2) It should also be used in combination with *d_inode()* to get the inode. (3) The dentry may have something attached to ->d_lower and the type field of the flags may be set to something other than miss or whiteout.

bool **d_really_is_positive**(const struct *dentry* *dentry)
Determine if a dentry is really positive (ignoring fallthroughs)

**Parameters**

**const struct dentry *dentry**
The dentry in question

**Description**

Returns true if the dentry represents a name that maps to an inode (ie. ->d_inode is not NULL). The dentry might still represent a whiteout if that is represented on medium as a 0,0 chardev.

Note! (1) This should be used *only* by a filesystem to examine its own dentries. It should not be used to look at some other filesystem's dentries. (2) It should also be used in combination with *d_inode()* to get the inode.

struct inode ***d_inode**(const struct *dentry* *dentry)

Get the actual inode of this dentry

**Parameters**

**const struct dentry *dentry**
The dentry to query

**Description**

This is the helper normal filesystems should use to get at their own inodes in their own dentries and ignore the layering superimposed upon them.

struct inode ***d_inode_rcu**(const struct *dentry* *dentry)

Get the actual inode of this dentry with READ_ONCE()

**Parameters**

**const struct dentry *dentry**
The dentry to query

**Description**

This is the helper normal filesystems should use to get at their own inodes in their own dentries and ignore the layering superimposed upon them.

struct inode ***d_backing_inode**(const struct dentry *upper)

Get upper or lower inode we should be using

**Parameters**

**const struct dentry *upper**
The upper layer

**Description**

This is the helper that should be used to get at the inode that will be used if this dentry were to be opened as a file. The inode may be on the upper dentry or it may be on a lower dentry pinned by the upper.

Normal filesystems should not use this to access their own inodes.

struct *dentry* ***d_real**(struct *dentry* *dentry, const struct *inode* *inode)

Return the real dentry

**Parameters**

**struct dentry *dentry**
the dentry to query

**const struct inode *inode**
inode to select the dentry from multiple layers (can be NULL)

**Description**

If dentry is on a union/overlay, then return the underlying, real dentry. Otherwise return the dentry itself.

See also: *Overview of the Linux Virtual File System*

struct inode *__d_real_inode__(const struct *dentry* *dentry)
Return the real inode

**Parameters**

**const struct dentry *dentry**
The dentry to query

**Description**

If dentry is on a union/overlay, then return the underlying, real inode. Otherwise return *d_inode()*.

### Inode Handling

int __inode_init_always__(struct super_block *sb, struct *inode* *inode)
perform inode structure initialisation

**Parameters**

**struct super_block *sb**
superblock inode belongs to

**struct inode *inode**
inode to initialise

**Description**

These are initializations that need to be done on every inode allocation as the fields are not initialised by slab allocation.

void __drop_nlink__(struct *inode* *inode)
directly drop an inode's link count

**Parameters**

**struct inode *inode**
inode

**Description**

This is a low-level filesystem helper to replace any direct filesystem manipulation of i_nlink. In cases where we are attempting to track writes to the filesystem, a decrement to zero means an imminent write when the file is truncated and actually unlinked on the filesystem.

void __clear_nlink__(struct *inode* *inode)
directly zero an inode's link count

**Parameters**

**struct inode *inode**
inode

**Description**

This is a low-level filesystem helper to replace any direct filesystem manipulation of i_nlink. See *drop_nlink()* for why we care about i_nlink hitting zero.

void **set_nlink**(struct *inode* *inode, unsigned int nlink)
directly set an inode's link count

**Parameters**

**struct inode *inode**
inode

**unsigned int nlink**
new nlink (should be non-zero)

**Description**

This is a low-level filesystem helper to replace any direct filesystem manipulation of i_nlink.

void **inc_nlink**(struct *inode* *inode)
directly increment an inode's link count

**Parameters**

**struct inode *inode**
inode

**Description**

This is a low-level filesystem helper to replace any direct filesystem manipulation of i_nlink. Currently, it is only here for parity with dec_nlink().

void **inode_sb_list_add**(struct *inode* *inode)
add inode to the superblock list of inodes

**Parameters**

**struct inode *inode**
inode to add

void **__insert_inode_hash**(struct *inode* *inode, unsigned long hashval)
hash an inode

**Parameters**

**struct inode *inode**
unhashed inode

**unsigned long hashval**
unsigned long value used to locate this object in the inode_hashtable.

Add an inode to the inode hash for this superblock.

void **__remove_inode_hash**(struct *inode* *inode)
remove an inode from the hash

**Parameters**

**struct inode *inode**
inode to unhash

Remove an inode from the superblock.

void **evict_inodes**(struct super_block *sb)
evict all evictable inodes for a superblock

**Parameters**

**struct super_block *sb**
superblock to operate on

**Description**

Make sure that no inodes with zero refcount are retained. This is called by superblock shutdown after having SB_ACTIVE flag removed, so any inode reaching zero refcount during or after that call will be immediately evicted.

struct inode *__new_inode__(struct super_block *sb)
obtain an inode

**Parameters**

**struct super_block *sb**
superblock

Allocates a new inode for given superblock. The default gfp_mask for allocations related to inode->i_mapping is GFP_HIGHUSER_MOVABLE. If HIGHMEM pages are unsuitable or it is known that pages allocated for the page cache are not reclaimable or migratable, mapping_set_gfp_mask() must be called with suitable flags on the newly created inode's mapping

void **unlock_new_inode**(struct *inode* *inode)
clear the I_NEW state and wake up any waiters

**Parameters**

**struct inode *inode**
new inode to unlock

**Description**

Called when the inode is fully initialised to clear the new state of the inode and wake up anyone waiting for the inode to finish initialisation.

void **lock_two_nondirectories**(struct inode *inode1, struct inode *inode2)
take two i_mutexes on non-directory objects

**Parameters**

**struct inode *inode1**
first inode to lock

**struct inode *inode2**
second inode to lock

**Description**

Lock any non-NULL argument. Passed objects must not be directories. Zero, one or two objects may be locked by this function.

void **unlock_two_nondirectories**(struct inode *inode1, struct inode *inode2)
release locks from *lock_two_nondirectories()*

**Parameters**

**struct inode *inode1**
first inode to unlock

**struct inode *inode2**
second inode to unlock

struct *inode* ***inode_insert5**(struct *inode* *inode, unsigned long hashval, int (*test)(struct *inode*, void*), int (*set)(struct *inode*, void*), void *data)

obtain an inode from a mounted file system

**Parameters**

**struct inode *inode**
pre-allocated inode to use for insert to cache

**unsigned long hashval**
hash value (usually inode number) to get

**int (*test)(struct inode *, void *)**
callback used for comparisons between inodes

**int (*set)(struct inode *, void *)**
callback used to initialize a new struct inode

**void *data**
opaque data pointer to pass to **test** and **set**

**Description**

Search for the inode specified by **hashval** and **data** in the inode cache, and if present it is return it with an increased reference count. This is a variant of *iget5_locked()* for callers that don't want to fail on memory allocation of inode.

If the inode is not in cache, insert the pre-allocated inode to cache and return it locked, hashed, and with the I_NEW flag set. The file system gets to fill it in before unlocking it via *unlock_new_inode()*.

Note both **test** and **set** are called with the inode_hash_lock held, so can't sleep.

struct inode ***iget5_locked**(struct super_block *sb, unsigned long hashval, int (*test)(struct inode*, void*), int (*set)(struct inode*, void*), void *data)

obtain an inode from a mounted file system

**Parameters**

**struct super_block *sb**
super block of file system

**unsigned long hashval**
hash value (usually inode number) to get

**int (*test)(struct inode *, void *)**
callback used for comparisons between inodes

**int (*set)(struct inode *, void *)**
callback used to initialize a new struct inode

**void *data**
opaque data pointer to pass to **test** and **set**

**Description**

Search for the inode specified by **hashval** and **data** in the inode cache, and if present it is return it with an increased reference count. This is a generalized version of *iget_locked()* for file systems where the inode number is not sufficient for unique identification of an inode.

If the inode is not in cache, allocate a new inode and return it locked, hashed, and with the I_NEW flag set. The file system gets to fill it in before unlocking it via *unlock_new_inode()*.

Note both **test** and **set** are called with the inode_hash_lock held, so can't sleep.

struct inode *__iget_locked__(struct super_block *sb, unsigned long ino)

obtain an inode from a mounted file system

**Parameters**

**struct super_block *sb**
super block of file system

**unsigned long ino**
inode number to get

**Description**

Search for the inode specified by **ino** in the inode cache and if present return it with an increased reference count. This is for file systems where the inode number is sufficient for unique identification of an inode.

If the inode is not in cache, allocate a new inode and return it locked, hashed, and with the I_NEW flag set. The file system gets to fill it in before unlocking it via *unlock_new_inode()*.

ino_t __iunique__(struct super_block *sb, ino_t max_reserved)

get a unique inode number

**Parameters**

**struct super_block *sb**
superblock

**ino_t max_reserved**
highest reserved inode number

Obtain an inode number that is unique on the system for a given superblock. This is used by file systems that have no natural permanent inode numbering system. An inode number is returned that is higher than the reserved limit but unique.

BUGS: With a large number of inodes live on the file system this function currently becomes quite slow.

struct inode *__ilookup5_nowait__(struct super_block *sb, unsigned long hashval, int (*test)(struct inode*, void*), void *data)

search for an inode in the inode cache

**Parameters**

**struct super_block *sb**
super block of file system to search

**unsigned long hashval**
hash value (usually inode number) to search for

**int (*test)(struct inode *, void *)**
callback used for comparisons between inodes

**void *data**
opaque data pointer to pass to **test**

**Description**

Search for the inode specified by **hashval** and **data** in the inode cache. If the inode is in the cache, the inode is returned with an incremented reference count.

Note2: **test** is called with the inode_hash_lock held, so can't sleep.

**Note**

I_NEW is not waited upon so you have to be very careful what you do with the returned inode. You probably should be using *ilookup5()* instead.

struct inode *__ilookup5__(struct super_block *sb, unsigned long hashval, int (*test)(struct inode*, void*), void *data)
search for an inode in the inode cache

**Parameters**

**struct super_block *sb**
super block of file system to search

**unsigned long hashval**
hash value (usually inode number) to search for

**int (*test)(struct inode *, void *)**
callback used for comparisons between inodes

**void *data**
opaque data pointer to pass to **test**

**Description**

Search for the inode specified by **hashval** and **data** in the inode cache, and if the inode is in the cache, return the inode with an incremented reference count. Waits on I_NEW before returning the inode. returned with an incremented reference count.

This is a generalized version of *ilookup()* for file systems where the inode number is not sufficient for unique identification of an inode.

**Note**

**test** is called with the inode_hash_lock held, so can't sleep.

struct inode *__ilookup__(struct super_block *sb, unsigned long ino)
search for an inode in the inode cache

**Parameters**

**struct super_block *sb**
super block of file system to search

**unsigned long ino**
inode number to search for

**Description**

Search for the inode **ino** in the inode cache, and if the inode is in the cache, the inode is returned with an incremented reference count.

struct inode *__find_inode_nowait__(struct super_block *sb, unsigned long hashval, int (*match)(struct inode*, unsigned long, void*), void *data)

find an inode in the inode cache

**Parameters**

**struct super_block *sb**
super block of file system to search

**unsigned long hashval**
hash value (usually inode number) to search for

**int (*match)(struct inode *, unsigned long, void *)**
callback used for comparisons between inodes

**void *data**
opaque data pointer to pass to **match**

**Description**

Search for the inode specified by **hashval** and **data** in the inode cache, where the helper function **match** will return 0 if the inode does not match, 1 if the inode does match, and -1 if the search should be stopped. The **match** function must be responsible for taking the i_lock spin_lock and checking i_state for an inode being freed or being initialized, and incrementing the reference count before returning 1. It also must not sleep, since it is called with the inode_hash_lock spinlock held.

This is a even more generalized version of *ilookup5()* when the function must never block --- find_inode() can block in __wait_on_freeing_inode() --- or when the caller can not increment the reference count because the resulting *iput()* might cause an inode eviction. The tradeoff is that the **match** funtion must be very carefully implemented.

struct inode *__find_inode_rcu__(struct super_block *sb, unsigned long hashval, int (*test)(struct inode*, void*), void *data)

find an inode in the inode cache

**Parameters**

**struct super_block *sb**
Super block of file system to search

**unsigned long hashval**
Key to hash

**int (*test)(struct inode *, void *)**
Function to test match on an inode

**void *data**
Data for test function

**Description**

Search for the inode specified by **hashval** and **data** in the inode cache, where the helper function **test** will return 0 if the inode does not match and 1 if it does. The **test** function must be responsible for taking the i_lock spin_lock and checking i_state for an inode being freed or being initialized.

If successful, this will return the inode for which the **test** function returned 1 and NULL otherwise.

The **test** function is not permitted to take a ref on any inode presented. It is also not permitted to sleep.

The caller must hold the RCU read lock.

struct inode *__find_inode_by_ino_rcu__(struct super_block *sb, unsigned long ino)

Find an inode in the inode cache

**Parameters**

**struct super_block *sb**
Super block of file system to search

**unsigned long ino**
The inode number to match

**Description**

Search for the inode specified by **hashval** and **data** in the inode cache, where the helper function **test** will return 0 if the inode does not match and 1 if it does. The **test** function must be responsible for taking the i_lock spin_lock and checking i_state for an inode being freed or being initialized.

If successful, this will return the inode for which the **test** function returned 1 and NULL otherwise.

The **test** function is not permitted to take a ref on any inode presented. It is also not permitted to sleep.

The caller must hold the RCU read lock.

void __iput__(struct *inode* *inode)

put an inode

**Parameters**

**struct inode *inode**
inode to put

Puts an inode, dropping its usage count. If the inode use count hits zero, the inode is then freed and may also be destroyed.

Consequently, *iput()* can sleep.

int __bmap__(struct *inode* *inode, sector_t *block)

find a block number in a file

**Parameters**

**struct inode *inode**
inode owning the block number being requested

**sector_t *block**
pointer containing the block to find

Replaces the value in *block with the block number on the device holding corresponding to the requested block number in the file. That is, asked for block 4 of inode 1 the function

will replace the 4 in *`block`, with disk block relative to the disk start that holds that block of the file.

Returns -EINVAL in case of error, 0 otherwise. If mapping falls into a hole, returns 0 and *`block` is also set to 0.

int **inode_update_timestamps**(struct *inode* *inode, int flags)

update the timestamps on the inode

**Parameters**

**struct inode *inode**
inode to be updated

**int flags**
S_* flags that needed to be updated

**Description**

The update_time function is called when an inode's timestamps need to be updated for a read or write operation. This function handles updating the actual timestamps. It's up to the caller to ensure that the inode is marked dirty appropriately.

In the case where any of S_MTIME, S_CTIME, or S_VERSION need to be updated, attempt to update all three of them. S_ATIME updates can be handled independently of the rest.

Returns a set of S_* flags indicating which values changed.

int **generic_update_time**(struct *inode* *inode, int flags)

update the timestamps on the inode

**Parameters**

**struct inode *inode**
inode to be updated

**int flags**
S_* flags that needed to be updated

**Description**

The update_time function is called when an inode's timestamps need to be updated for a read or write operation. In the case where any of S_MTIME, S_CTIME, or S_VERSION need to be updated we attempt to update all three of them. S_ATIME updates can be handled done independently of the rest.

Returns a S_* mask indicating which fields were updated.

int **file_remove_privs**(struct *file* *file)

remove special file privileges (suid, capabilities)

**Parameters**

**struct file *file**
file to remove privileges from

**Description**

When file is modified by a write or truncation ensure that special file privileges are removed.

**Return**

0 on success, negative errno on failure.

int **file_update_time**(struct *file* *file)
update mtime and ctime time

**Parameters**

**struct file *file**
file accessed

**Description**

Update the mtime and ctime members of an inode and mark the inode for writeback. Note that this function is meant exclusively for usage in the file write path of filesystems, and filesystems may choose to explicitly ignore updates via this function with the _NOCMTIME inode flag, e.g. for network filesystem where these imestamps are handled by the server. This can return an error for file systems who need to allocate space in order to update an inode.

**Return**

0 on success, negative errno on failure.

int **file_modified**(struct *file* *file)
handle mandated vfs changes when modifying a file

**Parameters**

**struct file *file**
file that was modified

**Description**

When file has been modified ensure that special file privileges are removed and time settings are updated.

**Context**

Caller must hold the file's inode lock.

**Return**

0 on success, negative errno on failure.

int **kiocb_modified**(struct kiocb *iocb)
handle mandated vfs changes when modifying a file

**Parameters**

**struct kiocb *iocb**
iocb that was modified

**Description**

When file has been modified ensure that special file privileges are removed and time settings are updated.

**Context**

Caller must hold the file's inode lock.

**Return**

0 on success, negative errno on failure.

void **inode_init_owner**(struct mnt_idmap *idmap, struct *inode* *inode, const struct *inode* *dir, umode_t mode)

Init uid,gid,mode for new inode according to posix standards

**Parameters**

**struct mnt_idmap *idmap**
idmap of the mount the inode was created from

**struct inode *inode**
New inode

**const struct inode *dir**
Directory inode

**umode_t mode**
mode of the new inode

**Description**

If the inode has been created through an idmapped mount the idmap of the vfsmount must be passed through **idmap**. This function will then take care to map the inode according to **idmap** before checking permissions and initializing i_uid and i_gid. On non-idmapped mounts or if permission checking is to be performed on the raw inode simply pass **nop_mnt_idmap**.

bool **inode_owner_or_capable**(struct mnt_idmap *idmap, const struct *inode* *inode)

check current task permissions to inode

**Parameters**

**struct mnt_idmap *idmap**
idmap of the mount the inode was found from

**const struct inode *inode**
inode being checked

**Description**

Return true if current either has CAP_FOWNER in a namespace with the inode owner uid mapped, or owns the file.

If the inode has been found through an idmapped mount the idmap of the vfsmount must be passed through **idmap**. This function will then take care to map the inode according to **idmap** before checking permissions. On non-idmapped mounts or if permission checking is to be performed on the raw inode simply pass **nop_mnt_idmap**.

void **inode_dio_wait**(struct *inode* *inode)

wait for outstanding DIO requests to finish

**Parameters**

**struct inode *inode**
inode to wait for

**Description**

Waits for all pending direct I/O requests to finish so that we can proceed with a truncate or equivalent operation.

Must be called under a lock that serializes taking new references to i_dio_count, usually by inode->i_mutex.

struct timespec64 **timestamp_truncate**(struct timespec64 t, struct *inode* *inode)

Truncate timespec to a granularity

**Parameters**

**struct timespec64 t**
Timespec

**struct inode *inode**
inode being updated

**Description**

Truncate a timespec to the granularity supported by the fs containing the inode. Always rounds down. gran must not be 0 nor greater than a second (NSEC_PER_SEC, or 10^9 ns).

struct timespec64 **current_time**(struct *inode* *inode)

Return FS time

**Parameters**

**struct inode *inode**
inode.

**Description**

Return the current time truncated to the time granularity supported by the fs.

Note that inode and inode->sb cannot be NULL. Otherwise, the function warns and returns time without truncation.

struct timespec64 **inode_set_ctime_current**(struct *inode* *inode)

set the ctime to current_time

**Parameters**

**struct inode *inode**
inode

**Description**

Set the inode->i_ctime to the current value for the inode. Returns the current value that was assigned to i_ctime.

umode_t **mode_strip_sgid**(struct mnt_idmap *idmap, const struct inode *dir, umode_t mode)

handle the sgid bit for non-directories

**Parameters**

**struct mnt_idmap *idmap**
idmap of the mount the inode was created from

**const struct inode *dir**
parent directory inode

**umode_t mode**
mode of the file to be created in **dir**

**Description**

If the **mode** of the new file has both the S_ISGID and S_IXGRP bit raised and **dir** has the S_ISGID bit raised ensure that the caller is either in the group of the parent directory or they

have CAP_FSETID in their user namespace and are privileged over the parent directory. In all other cases, strip the S_ISGID bit from **mode**.

**Return**

the new mode to use for the file

void **make_bad_inode**(struct *inode* *inode)
mark an inode bad due to an I/O error

**Parameters**

**struct inode *inode**
Inode to mark bad

When an inode cannot be read due to a media or remote network failure this function makes the inode ”bad” and causes I/O operations on it to fail from this point on.

bool **is_bad_inode**(struct *inode* *inode)
is an inode errored

**Parameters**

**struct inode *inode**
inode to test

Returns true if the inode in question has been marked as bad.

void **iget_failed**(struct *inode* *inode)
Mark an under-construction inode as dead and release it

**Parameters**

**struct inode *inode**
The inode to discard

**Description**

Mark an under-construction inode as dead and release it.

### Registration and Superblocks

void **deactivate_locked_super**(struct super_block *s)
drop an active reference to superblock

**Parameters**

**struct super_block *s**
superblock to deactivate

Drops an active reference to superblock, converting it into a temporary one if there is no other active references left. In that case we tell fs driver to shut it down and drop the temporary reference we had just acquired.

Caller holds exclusive lock on superblock; that lock is released.

void **deactivate_super**(struct super_block *s)
drop an active reference to superblock

**Parameters**

**struct super_block *s**
superblock to deactivate

Variant of *deactivate_locked_super()*, except that superblock is *not* locked by caller. If we are going to drop the final active reference, lock will be acquired prior to that.

void **retire_super**(struct super_block *sb)
prevents superblock from being reused

**Parameters**

**struct super_block *sb**
superblock to retire

The function marks superblock to be ignored in superblock test, which prevents it from being reused for any new mounts. If the superblock has a private bdi, it also unregisters it, but doesn't reduce the refcount of the superblock to prevent potential races. The refcount is reduced by *generic_shutdown_super()*. The function can not be called concurrently with *generic_shutdown_super()*. It is safe to call the function multiple times, subsequent calls have no effect.

The marker will affect the re-use only for block-device-based superblocks. Other superblocks will still get marked if this function is used, but that will not affect their reusability.

void **generic_shutdown_super**(struct super_block *sb)
common helper for ->kill_sb()

**Parameters**

**struct super_block *sb**
superblock to kill

*generic_shutdown_super()* does all fs-independent work on superblock shutdown. Typical ->kill_sb() should pick all fs-specific objects that need destruction out of superblock, call *generic_shutdown_super()* and release aforementioned objects. Note: dentries and inodes _are_ taken care of and do not need specific handling.

Upon calling this function, the filesystem may no longer alter or rearrange the set of dentries belonging to this super_block, nor may it change the attachments of dentries to inodes.

struct super_block *__sget_fc__(struct fs_context *fc, int (*test)(struct super_block*, struct fs_context*), int (*set)(struct super_block*, struct fs_context*))
Find or create a superblock

**Parameters**

**struct fs_context *fc**
Filesystem context.

**int (*test)(struct super_block *, struct fs_context *)**
Comparison callback

**int (*set)(struct super_block *, struct fs_context *)**
Setup callback

**Description**

Create a new superblock or find an existing one.

The **test** callback is used to find a matching existing superblock. Whether or not the requested parameters in **fc** are taken into account is specific to the **test** callback that is used. They may even be completely ignored.

If an extant superblock is matched, it will be returned unless:

(1) the namespace the filesystem context **fc** and the extant superblock's namespace differ

(2) the filesystem context **fc** has requested that reusing an extant superblock is not allowed

In both cases EBUSY will be returned.

If no match is made, a new superblock will be allocated and basic initialisation will be performed (s_type, s_fs_info and s_id will be set and the **set** callback will be invoked), the superblock will be published and it will be returned in a partially constructed state with SB_BORN and SB_ACTIVE as yet unset.

**Return**

**On success, an extant or newly created superblock is**
returned. On failure an error pointer is returned.

struct super_block *`sget`(struct file_system_type *type, int (*test)(struct super_block*, void*), int (*set)(struct super_block*, void*), int flags, void *data)

find or create a superblock

**Parameters**

**`struct file_system_type *type`**
filesystem type superblock should belong to

**`int (*test)(struct super_block *,void *)`**
comparison callback

**`int (*set)(struct super_block *,void *)`**
setup callback

**`int flags`**
mount flags

**`void *data`**
argument to each of them

void `iterate_supers_type`(struct file_system_type *type, void (*f)(struct super_block*, void*), void *arg)

call function for superblocks of given type

**Parameters**

**`struct file_system_type *type`**
fs type

**`void (*f)(struct super_block *, void *)`**
function to call

**`void *arg`**
argument to pass to it

Scans the superblock list and calls given function, passing it locked superblock and given argument.

int **get_anon_bdev**(dev_t *p)

Allocate a block device for filesystems which don't have one.

**Parameters**

**dev_t *p**
Pointer to a dev_t.

**Description**

Filesystems which don't use real block devices can call this function to allocate a virtual block device.

**Context**

Any context. Frequently called while holding sb_lock.

**Return**

0 on success, -EMFILE if there are no anonymous bdevs left or -ENOMEM if memory allocation failed.

struct super_block ***sget_dev**(struct fs_context *fc, dev_t dev)

Find or create a superblock by device number

**Parameters**

**struct fs_context *fc**
Filesystem context.

**dev_t dev**
device number

**Description**

Find or create a superblock using the provided device number that will be stored in fc->sget_key.

If an extant superblock is matched, then that will be returned with an elevated reference count that the caller must transfer or discard.

If no match is made, a new superblock will be allocated and basic initialisation will be performed (s_type, s_fs_info, s_id, s_dev will be set). The superblock will be published and it will be returned in a partially constructed state with SB_BORN and SB_ACTIVE as yet unset.

**Return**

**an existing or newly created superblock on success, an error**
pointer on failure.

int **get_tree_bdev**(struct fs_context *fc, int (*fill_super)(struct super_block*, struct fs_context*))

Get a superblock based on a single block device

**Parameters**

**struct fs_context *fc**
The filesystem context holding the parameters

**int (*fill_super)(struct super_block *, struct fs_context *)**
Helper to initialise a new superblock

int **vfs_get_tree**(struct fs_context *fc)

Get the mountable root

**Parameters**

**struct fs_context *fc**
The superblock configuration context.

**Description**

The filesystem is invoked to get or create a superblock which can then later be used for mounting. The filesystem places a pointer to the root to be used for mounting in **fc->root**.

int **freeze_super**(struct super_block *sb, enum *freeze_holder* who)

lock the filesystem and force it into a consistent state

**Parameters**

**struct super_block *sb**
the super to lock

**enum freeze_holder who**
context that wants to freeze

**Description**

Syncs the super to make sure the filesystem is consistent and calls the fs's freeze_fs. Subsequent calls to this without first thawing the fs may return -EBUSY.

**who** should be: * FREEZE_HOLDER_USERSPACE if userspace wants to freeze the fs; * FREEZE_HOLDER_KERNEL if the kernel wants to freeze the fs. * FREEZE_MAY_NEST whether nesting freeze and thaw requests is allowed.

The **who** argument distinguishes between the kernel and userspace trying to freeze the filesystem. Although there cannot be multiple kernel freezes or multiple userspace freezes in effect at any given time, the kernel and userspace can both hold a filesystem frozen. The filesystem remains frozen until there are no kernel or userspace freezes in effect.

A filesystem may hold multiple devices and thus a filesystems may be frozen through the block layer via multiple block devices. In this case the request is marked as being allowed to nest by passing FREEZE_MAY_NEST. The filesystem remains frozen until all block devices are unfrozen. If multiple freezes are attempted without FREEZE_MAY_NEST -EBUSY will be returned.

During this function, sb->s_writers.frozen goes through these values:

SB_UNFROZEN: File system is normal, all writes progress as usual.

SB_FREEZE_WRITE: The file system is in the process of being frozen. New writes should be blocked, though page faults are still allowed. We wait for all writes to complete and then proceed to the next stage.

SB_FREEZE_PAGEFAULT: Freezing continues. Now also page faults are blocked but internal fs threads can still modify the filesystem (although they should not dirty new pages or inodes), writeback can run etc. After waiting for all running page faults we sync the filesystem which will clean all dirty pages and inodes (no new dirty pages or inodes can be created when sync is running).

SB_FREEZE_FS: The file system is frozen. Now all internal sources of fs modification are blocked (e.g. XFS preallocation truncation on inode reclaim). This is usually implemented by blocking new transactions for filesystems that have them and need this additional guard.

After all internal writers are finished we call ->freeze_fs() to finish filesystem freezing. Then we transition to SB_FREEZE_COMPLETE state. This state is mostly auxiliary for filesystems to verify they do not modify frozen fs.

sb->s_writers.frozen is protected by sb->s_umount.

**Return**

**If the freeze was successful zero is returned. If the freeze**
failed a negative error code is returned.

int **thaw_super**(struct super_block *sb, enum *freeze_holder* who)

- unlock filesystem

**Parameters**

**struct super_block *sb**
the super to thaw

**enum freeze_holder who**
context that wants to freeze

**Description**

Unlocks the filesystem and marks it writeable again after *freeze_super()* if there are no remaining freezes on the filesystem.

**who** should be: * FREEZE_HOLDER_USERSPACE if userspace wants to thaw the fs; * FREEZE_HOLDER_KERNEL if the kernel wants to thaw the fs. * FREEZE_MAY_NEST whether nesting freeze and thaw requests is allowed

A filesystem may hold multiple devices and thus a filesystems may have been frozen through the block layer via multiple block devices. The filesystem remains frozen until all block devices are unfrozen.

## File Locks

bool **locks_owner_has_blockers**(struct file_lock_context *flctx, fl_owner_t owner)
Check for blocking lock requests

**Parameters**

**struct file_lock_context *flctx**
file lock context

**fl_owner_t owner**
lock owner

**Description**

**Return values:**
true: **owner** has at least one blocker false: **owner** has no blockers

int **locks_delete_block**(struct file_lock *waiter)
stop waiting for a file lock

**Parameters**

**struct file_lock *waiter**
the lock which was waiting

lockd/nfsd need to disconnect the lock while working on it.

int **posix_lock_file**(struct file *filp, struct file_lock *fl, struct file_lock *conflock)
Apply a POSIX-style lock to a file

**Parameters**

**struct file *filp**
The file to apply the lock to

**struct file_lock *fl**
The lock to be applied

**struct file_lock *conflock**
Place to return a copy of the conflicting lock, if found.

**Description**

Add a POSIX style lock to a file. We merge adjacent & overlapping locks whenever possible. POSIX locks are sorted by owner task, then by starting address

Note that if called with an FL_EXISTS argument, the caller may determine whether or not a lock was successfully freed by testing the return value for -ENOENT.

int **__break_lease**(struct *inode* *inode, unsigned int mode, unsigned int type)
revoke all outstanding leases on file

**Parameters**

**struct inode *inode**
the inode of the file to return

**unsigned int mode**
O_RDONLY: break only write leases; O_WRONLY or O_RDWR: break all leases

**unsigned int type**
FL_LEASE: break leases and delegations; FL_DELEG: break only delegations

break_lease (inlined for speed) has checked there already is at least some kind of lock (maybe a lease) on this file. Leases are broken on a call to open() or truncate(). This function can sleep unless you specified `O_NONBLOCK` to your open().

void **lease_get_mtime**(struct *inode* *inode, struct timespec64 *time)
update modified time of an inode with exclusive lease

**Parameters**

**struct inode *inode**
the inode

**struct timespec64 *time**
pointer to a timespec which contains the last modified time

**Description**

This is to force NFS clients to flush their caches for files with exclusive leases. The justification is that if someone has an exclusive lease, then they could be modifying it.

int **generic_setlease**(struct file *filp, int arg, struct file_lock **flp, void **priv)
sets a lease on an open file

**Parameters**

**struct file *filp**
file pointer

**int arg**
type of lease to obtain

**struct file_lock **flp**
input - file_lock to use, output - file_lock inserted

**void **priv**
private data for lm_setup (may be NULL if lm_setup doesn't require it)

The (input) flp->fl_lmops->lm_break function is required by break_lease().

int **vfs_setlease**(struct file *filp, int arg, struct file_lock **lease, void **priv)
sets a lease on an open file

**Parameters**

**struct file *filp**
file pointer

**int arg**
type of lease to obtain

**struct file_lock **lease**
file_lock to use when adding a lease

**void **priv**
private info for lm_setup when adding a lease (may be NULL if lm_setup doesn't require it)

**Description**

Call this to establish a lease on the file. The ”lease” argument is not used for F_UNLCK requests and may be NULL. For commands that set or alter an existing lease, the `(*lease)->fl_lmops->lm_break` operation must be set; if not, this function will return -ENOLCK (and generate a scary-looking stack trace).

The ”priv” pointer is passed directly to the lm_setup function as-is. It may be NULL if the lm_setup operation doesn't require it.

int **locks_lock_inode_wait**(struct *inode* *inode, struct file_lock *fl)
Apply a lock to an inode

**Parameters**

**struct inode *inode**
inode of the file to apply to

**struct file_lock *fl**
The lock to be applied

**Description**

Apply a POSIX or FLOCK style lock request to an inode.

int **vfs_test_lock**(struct file *filp, struct file_lock *fl)
test file byte range lock

**Parameters**

**struct file *filp**
The file to test lock for

**struct file_lock *fl**
The lock to test; also used to hold result

**Description**

Returns -ERRNO on failure. Indicates presence of conflicting lock by setting conf->fl_type to something other than F_UNLCK.

int **vfs_lock_file**(struct file *filp, unsigned int cmd, struct file_lock *fl, struct file_lock *conf)
file byte range lock

**Parameters**

**struct file *filp**
The file to apply the lock to

**unsigned int cmd**
type of locking operation (F_SETLK, F_GETLK, etc.)

**struct file_lock *fl**
The lock to be applied

**struct file_lock *conf**
Place to return a copy of the conflicting lock, if found.

**Description**

A caller that doesn't care about the conflicting lock may pass NULL as the final argument.

If the filesystem defines a private ->lock() method, then **conf** will be left unchanged; so a caller that cares should initialize it to some acceptable default.

To avoid blocking kernel daemons, such as lockd, that need to acquire POSIX locks, the ->lock() interface may return asynchronously, before the lock has been granted or denied by the underlying filesystem, if (and only if) lm_grant is set. Additionally EXPORT_OP_ASYNC_LOCK in export_operations flags need to be set.

Callers expecting ->lock() to return asynchronously will only use F_SETLK, not F_SETLKW; they will set FL_SLEEP if (and only if) the request is for a blocking lock. When ->lock() does return asynchronously, it must return FILE_LOCK_DEFERRED, and call ->lm_grant() when the lock request completes. If the request is for non-blocking lock the file system should return FILE_LOCK_DEFERRED then try to get the lock and call the callback routine with the result. If the request timed out the callback routine will return a nonzero return code and the file system should release the lock. The file system is also responsible to keep a corresponding posix lock when it grants a lock so the VFS can find out which locks are locally held and do the correct lock cleanup when required. The underlying filesystem must not drop the kernel lock or call ->lm_grant() before returning to the caller with a FILE_LOCK_DEFERRED return code.

int **vfs_cancel_lock**(struct file *filp, struct file_lock *fl)
file byte range unblock lock

**Parameters**

**struct file *filp**
The file to apply the unblock to

**struct file_lock *fl**
The lock to be unblocked

**Description**

Used by lock managers to cancel blocked requests

bool **vfs_inode_has_locks**(struct *inode* *inode)
are any file locks held on **inode**?

**Parameters**

**struct inode *inode**
inode to check for locks

**Description**

Return true if there are any FL_POSIX or FL_FLOCK locks currently set on **inode**.

int **posix_lock_inode_wait**(struct *inode* *inode, struct file_lock *fl)
Apply a POSIX-style lock to a file

**Parameters**

**struct inode *inode**
inode of file to which lock request should be applied

**struct file_lock *fl**
The lock to be applied

**Description**

Apply a POSIX style lock request to an inode.

int **fcntl_getlease**(struct file *filp)
Enquire what lease is currently active

**Parameters**

**struct file *filp**
the file

The value returned by this function will be one of (if no lease break is pending):

F_RDLCK to indicate a shared lease is held.

F_WRLCK to indicate an exclusive lease is held.

F_UNLCK to indicate no lease is held.

(if a lease break is pending):

**F_RDLCK to indicate an exclusive lease needs to be**
changed to a shared lease (or removed).

F_UNLCK to indicate the lease needs to be removed.

XXX: sfr & willy disagree over whether F_INPROGRESS should be returned to userspace.

int **check_conflicting_open**(struct file *filp, const int arg, int flags)

see if the given file points to an inode that has an existing open that would conflict with the desired lease.

**Parameters**

**struct file *filp**
file to check

**const int arg**
type of lease that we're trying to acquire

**int flags**
current lock flags

**Description**

Check to see if there's an existing open fd on this file that would conflict with the lease we're trying to set.

int **fcntl_setlease**(unsigned int fd, struct file *filp, int arg)

sets a lease on an open file

**Parameters**

**unsigned int fd**
open file descriptor

**struct file *filp**
file pointer

**int arg**
type of lease to obtain

Call this fcntl to establish a lease on the file. Note that you also need to call `F_SETSIG` to receive a signal when the lease is broken.

int **flock_lock_inode_wait**(struct *inode* *inode, struct file_lock *fl)

Apply a FLOCK-style lock to a file

**Parameters**

**struct inode *inode**
inode of the file to apply to

**struct file_lock *fl**
The lock to be applied

**Description**

Apply a FLOCK style lock request to an inode.

long **sys_flock**(unsigned int fd, unsigned int cmd)

- flock() system call.

**Parameters**

**unsigned int fd**
the file descriptor to lock.

**unsigned int cmd**
the type of lock to apply.

Apply a FL_FLOCK style lock to an open file descriptor. The **cmd** can be one of:

- LOCK_SH -- a shared lock.
- LOCK_EX -- an exclusive lock.
- LOCK_UN -- remove an existing lock.
- LOCK_MAND -- a 'mandatory' flock. (DEPRECATED)

LOCK_MAND support has been removed from the kernel.

pid_t **locks_translate_pid**(struct file_lock *fl, struct pid_namespace *ns)
translate a file_lock's fl_pid number into a namespace

**Parameters**

**struct file_lock *fl**
The file_lock who's fl_pid should be translated

**struct pid_namespace *ns**
The namespace into which the pid should be translated

**Description**

Used to translate a fl_pid into a namespace virtual pid number

### Other Functions

void **mpage_readahead**(struct readahead_control *rac, get_block_t get_block)
start reads against pages

**Parameters**

**struct readahead_control *rac**
Describes which pages to read.

**get_block_t get_block**
The filesystem's block mapper function.

**Description**

This function walks the pages and the blocks within each page, building and emitting large BIOs.

If anything unusual happens, such as:

- encountering a page which has buffers
- encountering a page which has a non-hole after a hole
- encountering a page with non-contiguous blocks

then this code just gives up and calls the buffer_head-based read function. It does handle a page which has holes at the end - that is a common case: the end-of-file on blocksize < PAGE_SIZE setups.

BH_Boundary explanation:

There is a problem. The mpage read code assembles several pages, gets all their disk mappings, and then submits them all. That's fine, but obtaining the disk mappings may require I/O. Reads of indirect blocks, for example.

So an mpage read of the first 16 blocks of an ext2 file will cause I/O to be submitted in the following order:

12 0 1 2 3 4 5 6 7 8 9 10 11 13 14 15 16

because the indirect block has to be read to get the mappings of blocks 13,14,15,16. Obviously, this impacts performance.

So what we do it to allow the filesystem's get_block() function to set BH_Boundary when it maps block 11. BH_Boundary says: mapping of the block after this one will require I/O against a block which is probably close to this one. So you should push what I/O you have currently accumulated.

This all causes the disk requests to be issued in the correct order.

int **mpage_writepages**(struct *address_space* *mapping, struct writeback_control *wbc, get_block_t get_block)

walk the list of dirty pages of the given address space & writepage() all of them

**Parameters**

**struct address_space *mapping**
address space structure to write

**struct writeback_control *wbc**
subtract the number of written pages from ***wbc->nr_to_write**

**get_block_t get_block**
the filesystem's block mapper function.

**Description**

This is a library function, which implements the writepages() address_space_operation.

int **generic_permission**(struct mnt_idmap *idmap, struct *inode* *inode, int mask)

check for access rights on a Posix-like filesystem

**Parameters**

**struct mnt_idmap *idmap**
idmap of the mount the inode was found from

**struct inode *inode**
inode to check access rights for

**int mask**
right to check for (`MAY_READ`, `MAY_WRITE`, `MAY_EXEC`, `MAY_NOT_BLOCK` ...)

**Description**

Used to check for read/write/execute permissions on a file. We use ”fsuid” for this, letting us set arbitrary permissions for filesystem access without changing the ”normal” uids which are used for other things.

generic_permission is rcu-walk aware. It returns -ECHILD in case an rcu-walk request cannot be satisfied (eg. requires blocking or too much complexity). It would then be called again in ref-walk mode.

If the inode has been found through an idmapped mount the idmap of the vfsmount must be passed through **idmap**. This function will then take care to map the inode according to **idmap** before checking permissions. On non-idmapped mounts or if permission checking is to be performed on the raw inode simply pass **nop_mnt_idmap**.

int **inode_permission**(struct mnt_idmap *idmap, struct *inode* *inode, int mask)

Check for access rights to a given inode

**Parameters**

**struct mnt_idmap *idmap**
idmap of the mount the inode was found from

**struct inode *inode**
Inode to check permission on

**int mask**
Right to check for (MAY_READ, MAY_WRITE, MAY_EXEC)

**Description**

Check for read/write/execute permissions on an inode. We use fs[ug]id for this, letting us set arbitrary permissions for filesystem access without changing the "normal" UIDs which are used for other things.

When checking for MAY_APPEND, MAY_WRITE must also be set in **mask**.

void **path_get**(const struct *path* *path)

get a reference to a path

**Parameters**

**const struct path *path**
path to get the reference to

**Description**

Given a path increment the reference count to the dentry and the vfsmount.

void **path_put**(const struct *path* *path)

put a reference to a path

**Parameters**

**const struct path *path**
path to put the reference to

**Description**

Given a path decrement the reference count to the dentry and the vfsmount.

int **vfs_path_parent_lookup**(struct *filename* *filename, unsigned int flags, struct path *parent, struct qstr *last, int *type, const struct path *root)

lookup a parent path relative to a dentry-vfsmount pair

**Parameters**

**struct filename *filename**
filename structure

**unsigned int flags**
lookup flags

**struct path *parent**
pointer to struct path to fill

**struct qstr *last**
last component

**int *type**
type of the last component

**const struct path *root**
pointer to struct path of the base directory

int **vfs_path_lookup**(struct *dentry* *dentry, struct vfsmount *mnt, const char *name, unsigned int flags, struct *path* *path)

lookup a file path relative to a dentry-vfsmount pair

**Parameters**

**struct dentry *dentry**
pointer to dentry of the base directory

**struct vfsmount *mnt**
pointer to vfs mount of the base directory

**const char *name**
pointer to file name

**unsigned int flags**
lookup flags

**struct path *path**
pointer to struct path to fill

struct dentry ***try_lookup_one_len**(const char *name, struct dentry *base, int len)

filesystem helper to lookup single pathname component

**Parameters**

**const char *name**
pathname component to lookup

**struct dentry *base**
base directory to lookup from

**int len**
maximum length **len** should be interpreted to

**Description**

Look up a dentry by name in the dcache, returning NULL if it does not currently exist. The function does not try to create a dentry.

Note that this routine is purely a helper for filesystem usage and should not be called by generic code.

The caller must hold base->i_mutex.

struct dentry *__lookup_one_len__(const char *name, struct dentry *base, int len)

filesystem helper to lookup single pathname component

**Parameters**

**const char *name**
pathname component to lookup

**struct dentry *base**
base directory to lookup from

**int len**
maximum length **len** should be interpreted to

**Description**

Note that this routine is purely a helper for filesystem usage and should not be called by generic code.

The caller must hold base->i_mutex.

struct dentry *__lookup_one__(struct mnt_idmap *idmap, const char *name, struct dentry *base, int len)

filesystem helper to lookup single pathname component

**Parameters**

**struct mnt_idmap *idmap**
idmap of the mount the lookup is performed from

**const char *name**
pathname component to lookup

**struct dentry *base**
base directory to lookup from

**int len**
maximum length **len** should be interpreted to

**Description**

Note that this routine is purely a helper for filesystem usage and should not be called by generic code.

The caller must hold base->i_mutex.

struct dentry *__lookup_one_unlocked__(struct mnt_idmap *idmap, const char *name, struct dentry *base, int len)

filesystem helper to lookup single pathname component

**Parameters**

**struct mnt_idmap *idmap**
idmap of the mount the lookup is performed from

**const char *name**
pathname component to lookup

**struct dentry *base**
base directory to lookup from

**int len**
maximum length **len** should be interpreted to

**Description**

Note that this routine is purely a helper for filesystem usage and should not be called by generic code.

Unlike lookup_one_len, it should be called without the parent i_mutex held, and will take the i_mutex itself if necessary.

struct dentry *__lookup_one_positive_unlocked__(struct mnt_idmap *idmap, const char *name, struct dentry *base, int len)
filesystem helper to lookup single pathname component

**Parameters**

**struct mnt_idmap *idmap**
idmap of the mount the lookup is performed from

**const char *name**
pathname component to lookup

**struct dentry *base**
base directory to lookup from

**int len**
maximum length **len** should be interpreted to

**Description**

This helper will yield ERR_PTR(-ENOENT) on negatives. The helper returns known positive or ERR_PTR(). This is what most of the users want.

Note that pinned negative with unlocked parent _can_ become positive at any time, so callers of *lookup_one_unlocked()* need to be very careful; pinned positives have >d_inode stable, so this one avoids such problems.

Note that this routine is purely a helper for filesystem usage and should not be called by generic code.

The helper should be called without i_mutex held.

struct dentry *__lookup_one_len_unlocked__(const char *name, struct dentry *base, int len)
filesystem helper to lookup single pathname component

**Parameters**

**const char *name**
pathname component to lookup

**struct dentry *base**
base directory to lookup from

**int len**
maximum length **len** should be interpreted to

**Description**

Note that this routine is purely a helper for filesystem usage and should not be called by generic code.

Unlike lookup_one_len, it should be called without the parent i_mutex held, and will take the i_mutex itself if necessary.

int **vfs_create**(struct mnt_idmap *idmap, struct inode *dir, struct *dentry* *dentry, umode_t mode, bool want_excl)

create new file

**Parameters**

**struct mnt_idmap *idmap**
idmap of the mount the inode was found from

**struct inode *dir**
inode of **dentry**

**struct dentry *dentry**
pointer to dentry of the base directory

**umode_t mode**
mode of the new file

**bool want_excl**
whether the file must not yet exist

**Description**

Create a new file.

If the inode has been found through an idmapped mount the idmap of the vfsmount must be passed through **idmap**. This function will then take care to map the inode according to **idmap** before checking permissions. On non-idmapped mounts or if permission checking is to be performed on the raw inode simply pass **nop_mnt_idmap**.

struct file *__kernel_tmpfile_open__(struct mnt_idmap *idmap, const struct path *parentpath, umode_t mode, int open_flag, const struct *cred* *cred)

open a tmpfile for kernel internal use

**Parameters**

**struct mnt_idmap *idmap**
idmap of the mount the inode was found from

**const struct path *parentpath**
path of the base directory

**umode_t mode**
mode of the new tmpfile

**int open_flag**
flags

**const struct cred *cred**
credentials for open

**Description**

Create and open a temporary file. The file is not accounted in nr_files, hence this is only for kernel internal use, and must not be installed into file tables or such.

int **vfs_mknod**(struct mnt_idmap *idmap, struct inode *dir, struct *dentry* *dentry, umode_t mode, dev_t dev)

create device node or file

**Parameters**

**struct mnt_idmap *idmap**
idmap of the mount the inode was found from

**struct inode *dir**
inode of **dentry**

**struct dentry *dentry**
pointer to dentry of the base directory

**umode_t mode**
mode of the new device node or file

**dev_t dev**
device number of device to create

**Description**

Create a device node or file.

If the inode has been found through an idmapped mount the idmap of the vfsmount must be passed through **idmap**. This function will then take care to map the inode according to **idmap** before checking permissions. On non-idmapped mounts or if permission checking is to be performed on the raw inode simply pass **nop_mnt_idmap**.

int **vfs_mkdir**(struct mnt_idmap *idmap, struct inode *dir, struct *dentry* *dentry, umode_t mode)

create directory

**Parameters**

**struct mnt_idmap *idmap**
idmap of the mount the inode was found from

**struct inode *dir**
inode of **dentry**

**struct dentry *dentry**
pointer to dentry of the base directory

**umode_t mode**
mode of the new directory

**Description**

Create a directory.

If the inode has been found through an idmapped mount the idmap of the vfsmount must be passed through **idmap**. This function will then take care to map the inode according to **idmap** before checking permissions. On non-idmapped mounts or if permission checking is to be performed on the raw inode simply pass **nop_mnt_idmap**.

int **vfs_rmdir**(struct mnt_idmap *idmap, struct inode *dir, struct *dentry* *dentry)

remove directory

**Parameters**

**struct mnt_idmap *idmap**
idmap of the mount the inode was found from

**struct inode *dir**
inode of **dentry**

**struct dentry *dentry**
pointer to dentry of the base directory

**Description**

Remove a directory.

If the inode has been found through an idmapped mount the idmap of the vfsmount must be passed through **idmap**. This function will then take care to map the inode according to **idmap** before checking permissions. On non-idmapped mounts or if permission checking is to be performed on the raw inode simply pass **nop_mnt_idmap**.

int **vfs_unlink**(struct mnt_idmap *idmap, struct inode *dir, struct *dentry* *dentry, struct inode **delegated_inode)

unlink a filesystem object

**Parameters**

**struct mnt_idmap *idmap**
idmap of the mount the inode was found from

**struct inode *dir**
parent directory

**struct dentry *dentry**
victim

**struct inode **delegated_inode**
returns victim inode, if the inode is delegated.

**Description**

The caller must hold dir->i_mutex.

If vfs_unlink discovers a delegation, it will return -EWOULDBLOCK and return a reference to the inode in delegated_inode. The caller should then break the delegation on that inode and retry. Because breaking a delegation may take a long time, the caller should drop dir->i_mutex before doing so.

Alternatively, a caller may pass NULL for delegated_inode. This may be appropriate for callers that expect the underlying filesystem not to be NFS exported.

If the inode has been found through an idmapped mount the idmap of the vfsmount must be passed through **idmap**. This function will then take care to map the inode according to **idmap** before checking permissions. On non-idmapped mounts or if permission checking is to be performed on the raw inode simply pass **nop_mnt_idmap**.

int **vfs_symlink**(struct mnt_idmap *idmap, struct inode *dir, struct *dentry* *dentry, const char *oldname)

create symlink

**Parameters**

**struct mnt_idmap *idmap**
idmap of the mount the inode was found from

**struct inode *dir**
inode of **dentry**

**struct dentry *dentry**
pointer to dentry of the base directory

**const char *oldname**
name of the file to link to

**Description**

Create a symlink.

If the inode has been found through an idmapped mount the idmap of the vfsmount must be passed through **idmap**. This function will then take care to map the inode according to **idmap** before checking permissions. On non-idmapped mounts or if permission checking is to be performed on the raw inode simply pass **nop_mnt_idmap**.

int **vfs_link**(struct dentry *old_dentry, struct mnt_idmap *idmap, struct inode *dir, struct dentry *new_dentry, struct inode **delegated_inode)

create a new link

**Parameters**

**struct dentry *old_dentry**
object to be linked

**struct mnt_idmap *idmap**
idmap of the mount

**struct inode *dir**
new parent

**struct dentry *new_dentry**
where to create the new link

**struct inode **delegated_inode**
returns inode needing a delegation break

**Description**

The caller must hold dir->i_mutex

If vfs_link discovers a delegation on the to-be-linked file in need of breaking, it will return -EWOULDBLOCK and return a reference to the inode in delegated_inode. The caller should then break the delegation and retry. Because breaking a delegation may take a long time, the caller should drop the i_mutex before doing so.

Alternatively, a caller may pass NULL for delegated_inode. This may be appropriate for callers that expect the underlying filesystem not to be NFS exported.

If the inode has been found through an idmapped mount the idmap of the vfsmount must be passed through **idmap**. This function will then take care to map the inode according to **idmap** before checking permissions. On non-idmapped mounts or if permission checking is to be performed on the raw inode simply pass **nop_mnt_idmap**.

int **vfs_rename**(struct *renamedata* *rd)

rename a filesystem object

**Parameters**

**struct renamedata *rd**
pointer to *struct renamedata* info

**Description**

The caller must hold multiple mutexes--see lock_rename()).

If vfs_rename discovers a delegation in need of breaking at either the source or destination, it will return -EWOULDBLOCK and return a reference to the inode in delegated_inode. The caller should then break the delegation and retry. Because breaking a delegation may take a long time, the caller should drop all locks before doing so.

Alternatively, a caller may pass NULL for delegated_inode. This may be appropriate for callers that expect the underlying filesystem not to be NFS exported.

The worst of all namespace operations - renaming directory. "Perverted" doesn't even start to describe it. Somebody in UCB had a heck of a trip... Problems:

a) we can get into loop creation.

b) race potential - two innocent renames can create a loop together. That's where 4.4BSD screws up. Current fix: serialization on sb->s_vfs_rename_mutex. We might be more accurate, but that's another story.

c) we may have to lock up to _four_ objects - parents and victim (if it exists), and source (if it's a non-directory or a subdirectory that moves to different parent). And that - after we got ->i_mutex on parents (until then we don't know whether the target exists). Solution: try to be smart with locking order for inodes. We rely on the fact that tree topology may change only under ->s_vfs_rename_mutex _and_ that parent of the object we move will be locked. Thus we can rank directories by the tree (ancestors first) and rank all non-directories after them. That works since everybody except rename does "lock parent, lookup, lock child" and rename is under ->s_vfs_rename_mutex. HOWEVER, it relies on the assumption that any object with ->lookup() has no more than 1 dentry. If "hybrid" objects will ever appear, we'd better make sure that there's no link(2) for them.

d) conversion from fhandle to dentry may come in the wrong moment - when we are removing the target. Solution: we will have to grab ->i_mutex in the fhandle_to_dentry code. [FIXME - current nfsfh.c relies on ->i_mutex on parents, which works but leads to some truly excessive locking].

int **vfs_readlink**(struct *dentry* *dentry, char __user *buffer, int buflen)

copy symlink body into userspace buffer

**Parameters**

**struct dentry *dentry**
dentry on which to get symbolic link

**char __user *buffer**
user memory pointer

**int buflen**
size of buffer

**Description**

Does not touch atime. That's up to the caller if necessary

Does not call security hook.

const char *__vfs_get_link__(struct *dentry* *dentry, struct delayed_call *done)
get symlink body

**Parameters**

**struct dentry *dentry**
dentry on which to get symbolic link

**struct delayed_call *done**
caller needs to free returned data with this

**Description**

Calls security hook and i_op->get_link() on the supplied inode.

It does not touch atime. That's up to the caller if necessary.

Does not work on "special" symlinks like /proc/$$/fd/N

int __sync_mapping_buffers__(struct *address_space* *mapping)
write out & wait upon a mapping's "associated" buffers

**Parameters**

**struct address_space *mapping**
the mapping which wants those buffers written

**Description**

Starts I/O against the buffers at mapping->i_private_list, and waits upon that I/O.

Basically, this is a convenience function for fsync(). **mapping** is a file or directory which needs those buffers to be written for a successful fsync().

int __generic_buffers_fsync_noflush__(struct *file* *file, loff_t start, loff_t end, bool datasync)
generic buffer fsync implementation for simple filesystems with no inode lock

**Parameters**

**struct file *file**
file to synchronize

**loff_t start**
start offset in bytes

**loff_t end**
end offset in bytes (inclusive)

**bool datasync**
only synchronize essential metadata if true

**Description**

This is a generic implementation of the fsync method for simple filesystems which track all non-inode metadata in the buffers list hanging off the address_space structure.

int **generic_buffers_fsync**(struct *file* *file, loff_t start, loff_t end, bool datasync)

generic buffer fsync implementation for simple filesystems with no inode lock

**Parameters**

**struct file *file**
file to synchronize

**loff_t start**
start offset in bytes

**loff_t end**
end offset in bytes (inclusive)

**bool datasync**
only synchronize essential metadata if true

**Description**

This is a generic implementation of the fsync method for simple filesystems which track all non-inode metadata in the buffers list hanging off the address_space structure. This also makes sure that a device cache flush operation is called at the end.

void **mark_buffer_dirty**(struct buffer_head *bh)

mark a buffer_head as needing writeout

**Parameters**

**struct buffer_head *bh**
the buffer_head to mark dirty

**Description**

*mark_buffer_dirty()* will set the dirty bit against the buffer, then set its backing page dirty, then tag the page as dirty in the page cache and then attach the address_space's inode to its superblock's dirty inode list.

*mark_buffer_dirty()* is atomic. It takes bh->b_folio->mapping->i_private_lock, i_pages lock and mapping->host->i_lock.

struct buffer_head *__bdev_getblk__(struct block_device *bdev, sector_t block, unsigned size, gfp_t gfp)

Get a buffer_head in a block device's buffer cache.

**Parameters**

**struct block_device *bdev**
The block device.

**sector_t block**
The block number.

**unsigned size**
The size of buffer_heads for this **bdev**.

**gfp_t gfp**
The memory allocation flags to use.

**Return**

The buffer head, or NULL if memory could not be allocated.

struct buffer_head *__bread_gfp(struct block_device *bdev, sector_t block, unsigned size, gfp_t gfp)

reads a specified block and returns the bh

**Parameters**

**struct block_device *bdev**
the block_device to read from

**sector_t block**
number of block

**unsigned size**
size (in bytes) to read

**gfp_t gfp**
page allocation flag

Reads a specified block, and returns buffer head that contains it. The page cache can be allocated from non-movable area not to prevent page migration if you set gfp to zero. It returns NULL if the block was unreadable.

void **block_invalidate_folio**(struct *folio* *folio, size_t offset, size_t length)

Invalidate part or all of a buffer-backed folio.

**Parameters**

**struct folio *folio**
The folio which is affected.

**size_t offset**
start of the range to invalidate

**size_t length**
length of the range to invalidate

**Description**

*block_invalidate_folio()* is called when all or part of the folio has been invalidated by a truncate operation.

*block_invalidate_folio()* does not have to release all buffers, but it must ensure that no dirty buffer is left outside **offset** and that no I/O is underway against any of the blocks which are outside the truncation point. Because the caller is about to free (and possibly reuse) those blocks on-disk.

void **clean_bdev_aliases**(struct block_device *bdev, sector_t block, sector_t len)

clean a range of buffers in block device

**Parameters**

**struct block_device *bdev**
Block device to clean buffers in

**sector_t block**
Start of a range of blocks to clean

**sector_t len**
Number of blocks to clean

### Description

We are taking a range of blocks for data and we don't want writeback of any buffer-cache aliases starting from return from this function and until the moment when something will explicitly mark the buffer dirty (hopefully that will not happen until we will free that block ;-) We don't even need to mark it not-uptodate - nobody can expect anything from a newly allocated buffer anyway. We used to use unmap_buffer() for such invalidation, but that was wrong. We definitely don't want to mark the alias unmapped, for example - it would confuse anyone who might pick it with bread() afterwards...

Also.. Note that bforget() doesn't lock the buffer. So there can be writeout I/O going on against recently-freed buffers. We don't wait on that I/O in bforget() - it's more efficient to wait on the I/O only if we really need to. That happens here.

int **bh_uptodate_or_lock**(struct buffer_head *bh)

Test whether the buffer is uptodate

### Parameters

**struct buffer_head *bh**
struct buffer_head

### Description

Return true if the buffer is up-to-date and false, with the buffer locked, if not.

int **__bh_read**(struct buffer_head *bh, blk_opf_t op_flags, bool wait)

Submit read for a locked buffer

### Parameters

**struct buffer_head *bh**
struct buffer_head

**blk_opf_t op_flags**
appending REQ_OP_* flags besides REQ_OP_READ

**bool wait**
wait until reading finish

### Description

Returns zero on success or don't wait, and -EIO on error.

void **__bh_read_batch**(int nr, struct buffer_head *bhs[], blk_opf_t op_flags, bool force_lock)

Submit read for a batch of unlocked buffers

### Parameters

**int nr**
entry number of the buffer batch

**struct buffer_head *bhs[]**
a batch of struct buffer_head

**blk_opf_t op_flags**
appending REQ_OP_* flags besides REQ_OP_READ

**bool force_lock**
force to get a lock on the buffer if set, otherwise drops any buffer that cannot lock.

**Description**

Returns zero on success or don't wait, and -EIO on error.

void **bio_reset**(struct *bio* *bio, struct block_device *bdev, blk_opf_t opf)
reinitialize a bio

**Parameters**

**struct bio *bio**
bio to reset

**struct block_device *bdev**
block device to use the bio for

**blk_opf_t opf**
operation and flags for bio

**Description**

After calling *bio_reset()*, **bio** will be in the same state as a freshly allocated bio returned bio *bio_alloc_bioset()* - the only fields that are preserved are the ones that are initialized by *bio_alloc_bioset()*. See comment in struct bio.

void **bio_chain**(struct *bio* *bio, struct *bio* *parent)
chain bio completions

**Parameters**

**struct bio *bio**
the target bio

**struct bio *parent**
the parent bio of **bio**

**Description**

The caller won't have a bi_end_io called when **bio** completes - instead, **parent**'s bi_end_io won't be called until both **parent** and **bio** have completed; the chained bio will also be freed when it completes.

The caller must not set bi_private or bi_end_io in **bio**.

struct bio *__bio_alloc_bioset__(struct block_device *bdev, unsigned short nr_vecs, blk_opf_t opf, gfp_t gfp_mask, struct bio_set *bs)
allocate a bio for I/O

**Parameters**

**struct block_device *bdev**
block device to allocate the bio for (can be `NULL`)

**unsigned short nr_vecs**
number of bvecs to pre-allocate

**blk_opf_t opf**
operation and flags for bio

**gfp_t gfp_mask**
the GFP_* mask given to the slab allocator

**struct bio_set *bs**
: the bio_set to allocate from.

**Description**

Allocate a bio from the mempools in **bs**.

If __GFP_DIRECT_RECLAIM is set then bio_alloc will always be able to allocate a bio. This is due to the mempool guarantees. To make this work, callers must never allocate more than 1 bio at a time from the general pool. Callers that need to allocate more than 1 bio must always submit the previously allocated bio for IO before attempting to allocate a new one. Failure to do so can cause deadlocks under memory pressure.

Note that when running under submit_bio_noacct() (i.e. any block driver), bios are not submitted until after you return - see the code in submit_bio_noacct() that converts recursion into iteration, to prevent stack overflows.

This would normally mean allocating multiple bios under submit_bio_noacct() would be susceptible to deadlocks, but we have deadlock avoidance code that resubmits any blocked bios from a rescuer thread.

However, we do not guarantee forward progress for allocations from other mempools. Doing multiple allocations from the same mempool under submit_bio_noacct() should be avoided - instead, use bio_set's front_pad for per bio allocations.

**Return**

Pointer to new bio on success, NULL on failure.

struct bio ***bio_kmalloc**(unsigned short nr_vecs, gfp_t gfp_mask)
: kmalloc a bio

**Parameters**

**unsigned short nr_vecs**
: number of bio_vecs to allocate

**gfp_t gfp_mask**
: the GFP_* mask given to the slab allocator

**Description**

Use kmalloc to allocate a bio (including bvecs). The bio must be initialized using bio_init() before use. To free a bio returned from this function use kfree() after calling bio_uninit(). A bio returned from this function can be reused by calling bio_uninit() before calling bio_init() again.

Note that unlike bio_alloc() or *bio_alloc_bioset()* allocations from this function are not backed by a mempool can fail. Do not use this function for allocations in the file system I/O path.

**Return**

Pointer to new bio on success, NULL on failure.

void **bio_put**(struct *bio* *bio)
: release a reference to a bio

**Parameters**

**struct bio *bio**
: bio to release reference to

**Description**

Put a reference to a `struct bio`, either one you have gotten with bio_alloc, bio_get or bio_clone_*. The last put of a bio will free it.

struct bio *__bio_alloc_clone__(struct block_device *bdev, struct bio *bio_src, gfp_t gfp, struct bio_set *bs)

clone a bio that shares the original bio's biovec

**Parameters**

**struct block_device *bdev**
block_device to clone onto

**struct bio *bio_src**
bio to clone from

**gfp_t gfp**
allocation priority

**struct bio_set *bs**
bio_set to allocate from

**Description**

Allocate a new bio that is a clone of **bio_src**. The caller owns the returned bio, but not the actual data it points to.

The caller must ensure that the return bio is not freed before **bio_src**.

int __bio_init_clone__(struct block_device *bdev, struct *bio* *bio, struct *bio* *bio_src, gfp_t gfp)

clone a bio that shares the original bio's biovec

**Parameters**

**struct block_device *bdev**
block_device to clone onto

**struct bio *bio**
bio to clone into

**struct bio *bio_src**
bio to clone from

**gfp_t gfp**
allocation priority

**Description**

Initialize a new bio in caller provided memory that is a clone of **bio_src**. The caller owns the returned bio, but not the actual data it points to.

The caller must ensure that **bio_src** is not freed before **bio**.

int __bio_add_pc_page__(struct request_queue *q, struct *bio* *bio, struct *page* *page, unsigned int len, unsigned int offset)

attempt to add page to passthrough bio

**Parameters**

**struct request_queue *q**
the target queue

**struct bio *bio**
destination bio

**struct page *page**
page to add

**unsigned int len**
vec entry length

**unsigned int offset**
vec entry offset

**Description**

Attempt to add a page to the bio_vec maplist. This can fail for a number of reasons, such as the bio being full or target block device limitations. The target block device must allow bio's up to PAGE_SIZE, so it is always possible to add a single page to an empty bio.

This should only be used by passthrough bios.

int **bio_add_zone_append_page**(struct *bio* *bio, struct *page* *page, unsigned int len, unsigned int offset)

attempt to add page to zone-append bio

**Parameters**

**struct bio *bio**
destination bio

**struct page *page**
page to add

**unsigned int len**
vec entry length

**unsigned int offset**
vec entry offset

**Description**

Attempt to add a page to the bio_vec maplist of a bio that will be submitted for a zone-append request. This can fail for a number of reasons, such as the bio being full or the target block device is not a zoned block device or other limitations of the target block device. The target block device must allow bio's up to PAGE_SIZE, so it is always possible to add a single page to an empty bio.

**Return**

number of bytes added to the bio, or 0 in case of a failure.

void **__bio_add_page**(struct *bio* *bio, struct *page* *page, unsigned int len, unsigned int off)

add page(s) to a bio in a new segment

**Parameters**

**struct bio *bio**
destination bio

**struct page *page**
start page to add

**unsigned int len**
: length of the data to add, may cross pages

**unsigned int off**
: offset of the data relative to **page**, may cross pages

**Description**

Add the data at **page** + **off** to **bio** as a new bvec. The caller must ensure that **bio** has space for another bvec.

int **bio_add_page**(struct *bio* *bio, struct *page* *page, unsigned int len, unsigned int offset)
: attempt to add page(s) to bio

**Parameters**

**struct bio *bio**
: destination bio

**struct page *page**
: start page to add

**unsigned int len**
: vec entry length, may cross pages

**unsigned int offset**
: vec entry offset relative to **page**, may cross pages

  Attempt to add page(s) to the bio_vec maplist. This will only fail if either bio->bi_vcnt == bio->bi_max_vecs or it's a cloned bio.

bool **bio_add_folio**(struct *bio* *bio, struct *folio* *folio, size_t len, size_t off)
: Attempt to add part of a folio to a bio.

**Parameters**

**struct bio *bio**
: BIO to add to.

**struct folio *folio**
: Folio to add.

**size_t len**
: How many bytes from the folio to add.

**size_t off**
: First byte in this folio to add.

**Description**

Filesystems that use folios can call this function instead of calling *bio_add_page()* for each page in the folio. If **off** is bigger than PAGE_SIZE, this function can create a bio_vec that starts in a page after the bv_page. BIOs do not support folios that are 4GiB or larger.

**Return**

Whether the addition was successful.

int **bio_iov_iter_get_pages**(struct *bio* *bio, struct iov_iter *iter)
: add user or kernel pages to a bio

**Parameters**

**struct bio *bio**
bio to add pages to

**struct iov_iter *iter**
iov iterator describing the region to be added

**Description**

This takes either an iterator pointing to user memory, or one pointing to kernel pages (BVEC iterator). If we're adding user pages, we pin them and map them into the kernel. On IO completion, the caller should put those pages. For bvec based iterators *bio_iov_iter_get_pages()* uses the provided bvecs rather than copying them. Hence anyone issuing kiocb based IO needs to ensure the bvecs and pages stay referenced until the submitted I/O is completed by a call to ->ki_complete() or returns with an error other than -EIOCBQUEUED. The caller needs to check if the bio is flagged BIO_NO_PAGE_REF on IO completion. If it isn't, then pages should be released.

The function tries, but does not guarantee, to pin as many pages as fit into the bio, or are requested in **iter**, whatever is smaller. If MM encounters an error pinning the requested pages, it stops. Error is returned only if 0 pages could be pinned.

int **submit_bio_wait**(struct *bio* *bio)
submit a bio, and wait until it completes

**Parameters**

**struct bio *bio**
The `struct bio` which describes the I/O

**Description**

Simple wrapper around submit_bio(). Returns 0 on success, or the error from *bio_endio()* on failure.

WARNING: Unlike to how submit_bio() is usually used, this function does not result in bio reference to be consumed. The caller must drop the reference on his own.

void **bio_copy_data**(struct bio *dst, struct bio *src)
copy contents of data buffers from one bio to another

**Parameters**

**struct bio *dst**
destination bio

**struct bio *src**
source bio

**Description**

Stops when it reaches the end of either **src** or **dst** - that is, copies min(src->bi_size, dst->bi_size) bytes (or the equivalent for lists of bios).

void **bio_endio**(struct *bio* *bio)
end I/O on a bio

**Parameters**

**struct bio *bio**
bio

**Description**

*bio_endio()* will end I/O on the whole bio. *bio_endio()* is the preferred way to end I/O on a bio. No one should call bi_end_io() directly on a bio unless they own it and thus know that it has an end_io function.

*bio_endio()* can be called several times on a bio that has been chained using bio_chain(). The ->bi_end_io() function will only be called the last time.

struct *bio* ***bio_split**(struct *bio* *bio, int sectors, gfp_t gfp, struct bio_set *bs)

split a bio

**Parameters**

**struct bio *bio**
bio to split

**int sectors**
number of sectors to split from the front of **bio**

**gfp_t gfp**
gfp mask

**struct bio_set *bs**
bio set to allocate from

**Description**

Allocates and returns a new bio which represents **sectors** from the start of **bio**, and updates **bio** to represent the remaining sectors.

Unless this is a discard request the newly allocated bio will point to **bio**'s bi_io_vec. It is the caller's responsibility to ensure that neither **bio** nor **bs** are freed before the split bio.

void **bio_trim**(struct *bio* *bio, sector_t offset, sector_t size)

trim a bio

**Parameters**

**struct bio *bio**
bio to trim

**sector_t offset**
number of sectors to trim from the front of **bio**

**sector_t size**
size we want to trim **bio** to, in sectors

**Description**

This function is typically used for bios that are cloned and submitted to the underlying device in parts.

int **bioset_init**(struct bio_set *bs, unsigned int pool_size, unsigned int front_pad, int flags)

Initialize a bio_set

**Parameters**

**struct bio_set *bs**
pool to initialize

**unsigned int pool_size**
Number of bio and bio_vecs to cache in the mempool

**unsigned int front_pad**
Number of bytes to allocate in front of the returned bio

**int flags**
Flags to modify behavior, currently `BIOSET_NEED_BVECS` and `BIOSET_NEED_RESCUER`

**Description**

Set up a bio_set to be used with **bio_alloc_bioset**. Allows the caller to ask for a number of bytes to be allocated in front of the bio. Front pad allocation is useful for embedding the bio inside another structure, to avoid allocating extra data to go with the bio. Note that the bio must be embedded at the END of that structure always, or things will break badly. If `BIOSET_NEED_BVECS` is set in **flags**, a separate pool will be allocated for allocating iovecs. This pool is not needed e.g. for *bio_init_clone()*. If `BIOSET_NEED_RESCUER` is set, a workqueue is created which can be used to dispatch queued requests when the mempool runs out of space.

int **seq_open**(struct *file* *file, const struct seq_operations *op)
initialize sequential file

**Parameters**

**struct file *file**
file we initialize

**const struct seq_operations *op**
method table describing the sequence

*seq_open()* sets **file**, associating it with a sequence described by **op**. **op->start()** sets the iterator up and returns the first element of sequence. **op->stop()** shuts it down. **op->next()** returns the next element of sequence. **op->show()** prints element into the buffer. In case of error ->start() and ->next() return ERR_PTR(error). In the end of sequence they return `NULL`. ->show() returns 0 in case of success and negative number in case of error. Returning SEQ_SKIP means ”discard this element and move on”.

**Note**

**seq_open() will allocate a struct seq_file and store its**
pointer in **file->private_data**. This pointer should not be modified.

ssize_t **seq_read**(struct *file* *file, char __user *buf, size_t size, loff_t *ppos)
->read() method for sequential files.

**Parameters**

**struct file *file**
the file to read from

**char __user *buf**
the buffer to read to

**size_t size**
the maximum number of bytes to read

**loff_t *ppos**
the current position in the file

Ready-made ->f_op->read()

loff_t **seq_lseek**(struct *file* *file, loff_t offset, int whence)

->llseek() method for sequential files.

**Parameters**

**struct file *file**
the file in question

**loff_t offset**
new position

**int whence**
0 for absolute, 1 for relative position

Ready-made ->f_op->llseek()

int **seq_release**(struct *inode* *inode, struct *file* *file)

free the structures associated with sequential file.

**Parameters**

**struct inode *inode**
its inode

Frees the structures associated with sequential file; can be used as ->f_op->release() if you don't have private data to destroy.

**struct file *file**
file in question

void **seq_escape_mem**(struct seq_file *m, const char *src, size_t len, unsigned int flags, const char *esc)

print data into buffer, escaping some characters

**Parameters**

**struct seq_file *m**
target buffer

**const char *src**
source buffer

**size_t len**
size of source buffer

**unsigned int flags**
flags to pass to string_escape_mem()

**const char *esc**
set of characters that need escaping

**Description**

Puts data into buffer, replacing each occurrence of character from given class (defined by **flags** and **esc**) with printable escaped sequence.

Use seq_has_overflowed() to check for errors.

char *__mangle_path__(char *s, const char *p, const char *esc)

mangle and copy path to buffer beginning

**Parameters**

**char *s**
buffer start

**const char *p**
beginning of path in above buffer

**const char *esc**
set of characters that need escaping

Copy the path from **p** to **s**, replacing each occurrence of character from **esc** with usual octal escape. Returns pointer past last written character in **s**, or NULL in case of failure.

int __seq_path__(struct seq_file *m, const struct *path* *path, const char *esc)

seq_file interface to print a pathname

**Parameters**

**struct seq_file *m**
the seq_file handle

**const struct path *path**
the struct path to print

**const char *esc**
set of characters to escape in the output

**Description**

return the absolute path of 'path', as represented by the dentry / mnt pair in the path parameter.

int __seq_file_path__(struct seq_file *m, struct *file* *file, const char *esc)

seq_file interface to print a pathname of a file

**Parameters**

**struct seq_file *m**
the seq_file handle

**struct file *file**
the struct file to print

**const char *esc**
set of characters to escape in the output

**Description**

return the absolute path to the file.

int __seq_write__(struct seq_file *seq, const void *data, size_t len)

write arbitrary data to buffer

**Parameters**

**struct seq_file *seq**
seq_file identifying the buffer to which data should be written

**const void *data**
data address

**size_t len**
number of bytes

**Description**

Return 0 on success, non-zero otherwise.

void **seq_pad**(struct seq_file *m, char c)
write padding spaces to buffer

**Parameters**

**struct seq_file *m**
seq_file identifying the buffer to which data should be written

**char c**
the byte to append after padding if non-zero

struct hlist_node ***seq_hlist_start**(struct hlist_head *head, loff_t pos)
start an iteration of a hlist

**Parameters**

**struct hlist_head *head**
the head of the hlist

**loff_t pos**
the start position of the sequence

**Description**

Called at seq_file->op->start().

struct hlist_node ***seq_hlist_start_head**(struct hlist_head *head, loff_t pos)
start an iteration of a hlist

**Parameters**

**struct hlist_head *head**
the head of the hlist

**loff_t pos**
the start position of the sequence

**Description**

Called at seq_file->op->start(). Call this function if you want to print a header at the top of the output.

struct hlist_node ***seq_hlist_next**(void *v, struct hlist_head *head, loff_t *ppos)
move to the next position of the hlist

**Parameters**

**void *v**
the current iterator

**struct hlist_head *head**
the head of the hlist

**loff_t *ppos**
the current position

**Description**

Called at seq_file->op->next().

struct hlist_node ***seq_hlist_start_rcu**(struct hlist_head *head, loff_t pos)
start an iteration of a hlist protected by RCU

**Parameters**

**struct hlist_head *head**
the head of the hlist

**loff_t pos**
the start position of the sequence

**Description**

Called at seq_file->op->start().

This list-traversal primitive may safely run concurrently with the _rcu list-mutation primitives such as hlist_add_head_rcu() as long as the traversal is guarded by rcu_read_lock().

struct hlist_node ***seq_hlist_start_head_rcu**(struct hlist_head *head, loff_t pos)
start an iteration of a hlist protected by RCU

**Parameters**

**struct hlist_head *head**
the head of the hlist

**loff_t pos**
the start position of the sequence

**Description**

Called at seq_file->op->start(). Call this function if you want to print a header at the top of the output.

This list-traversal primitive may safely run concurrently with the _rcu list-mutation primitives such as hlist_add_head_rcu() as long as the traversal is guarded by rcu_read_lock().

struct hlist_node ***seq_hlist_next_rcu**(void *v, struct hlist_head *head, loff_t *ppos)
move to the next position of the hlist protected by RCU

**Parameters**

**void *v**
the current iterator

**struct hlist_head *head**
the head of the hlist

**loff_t *ppos**
the current position

**Description**

Called at seq_file->op->next().

This list-traversal primitive may safely run concurrently with the _rcu list-mutation primitives such as hlist_add_head_rcu() as long as the traversal is guarded by rcu_read_lock().

struct hlist_node *__seq_hlist_start_percpu__(struct hlist_head __percpu *head, int *cpu, loff_t pos)

start an iteration of a percpu hlist array

**Parameters**

**struct hlist_head __percpu *head**
pointer to percpu array of struct hlist_heads

**int *cpu**
pointer to cpu ”cursor”

**loff_t pos**
start position of sequence

**Description**

Called at seq_file->op->start().

struct hlist_node *__seq_hlist_next_percpu__(void *v, struct hlist_head __percpu *head, int *cpu, loff_t *pos)

move to the next position of the percpu hlist array

**Parameters**

**void *v**
pointer to current hlist_node

**struct hlist_head __percpu *head**
pointer to percpu array of struct hlist_heads

**int *cpu**
pointer to cpu ”cursor”

**loff_t *pos**
start position of sequence

**Description**

Called at seq_file->op->next().

int __register_filesystem__(struct file_system_type *fs)

register a new filesystem

**Parameters**

**struct file_system_type * fs**
the file system structure

Adds the file system passed to the list of file systems the kernel is aware of for mount and other syscalls. Returns 0 on success, or a negative errno code on an error.

The `struct file_system_type` that is passed is linked into the kernel structures and must not be freed until the file system has been unregistered.

int __unregister_filesystem__(struct file_system_type *fs)

unregister a file system

**Parameters**

**struct file_system_type * fs**
filesystem to unregister

Remove a file system that was previously successfully registered with the kernel. An error is returned if the file system is not found. Zero is returned on a success.

Once this function has returned the `struct file_system_type` structure may be freed or reused.

void **wbc_attach_and_unlock_inode**(struct writeback_control *wbc, struct *inode* *inode)
associate wbc with target inode and unlock it

**Parameters**

**struct writeback_control *wbc**
writeback_control of interest

**struct inode *inode**
target inode

**Description**

**inode** is locked and about to be written back under the control of **wbc**. Record **inode**'s writeback context into **wbc** and unlock the i_lock. On writeback completion, *wbc_detach_inode()* should be called. This is used to track the cgroup writeback context.

void **wbc_detach_inode**(struct writeback_control *wbc)
disassociate wbc from inode and perform foreign detection

**Parameters**

**struct writeback_control *wbc**
writeback_control of the just finished writeback

**Description**

To be called after a writeback attempt of an inode finishes and undoes *wbc_attach_and_unlock_inode()*. Can be called under any context.

As concurrent write sharing of an inode is expected to be very rare and memcg only tracks page ownership on first-use basis severely confining the usefulness of such sharing, cgroup writeback tracks ownership per-inode. While the support for concurrent write sharing of an inode is deemed unnecessary, an inode being written to by different cgroups at different points in time is a lot more common, and, more importantly, charging only by first-use can too readily lead to grossly incorrect behaviors (single foreign page can lead to gigabytes of writeback to be incorrectly attributed).

To resolve this issue, cgroup writeback detects the majority dirtier of an inode and transfers the ownership to it. To avoid unnecessary oscillation, the detection mechanism keeps track of history and gives out the switch verdict only if the foreign usage pattern is stable over a certain amount of time and/or writeback attempts.

On each writeback attempt, **wbc** tries to detect the majority writer using Boyer-Moore majority vote algorithm. In addition to the byte count from the majority voting, it also counts the bytes written for the current wb and the last round's winner wb (max of last round's current wb, the winner from two rounds ago, and the last round's majority candidate). Keeping track of the

historical winner helps the algorithm to semi-reliably detect the most active writer even when it's not the absolute majority.

Once the winner of the round is determined, whether the winner is foreign or not and how much IO time the round consumed is recorded in inode->i_wb_frn_history. If the amount of recorded foreign IO time is over a certain threshold, the switch verdict is given.

void **wbc_account_cgroup_owner**(struct writeback_control *wbc, struct *page* *page, size_t bytes)

account writeback to update inode cgroup ownership

**Parameters**

**struct writeback_control *wbc**
writeback_control of the writeback in progress

**struct page *page**
page being written out

**size_t bytes**
number of bytes being written out

**Description**

**bytes** from **page** are about to written out during the writeback controlled by **wbc**. Keep the book for foreign inode detection. See *wbc_detach_inode()*.

void **__mark_inode_dirty**(struct *inode* *inode, int flags)

internal function to mark an inode dirty

**Parameters**

**struct inode *inode**
inode to mark

**int flags**
what kind of dirty, e.g. I_DIRTY_SYNC. This can be a combination of multiple I_DIRTY_* flags, except that I_DIRTY_TIME can't be combined with I_DIRTY_PAGES.

**Description**

Mark an inode as dirty. We notify the filesystem, then update the inode's dirty flags. Then, if needed we add the inode to the appropriate dirty list.

Most callers should use mark_inode_dirty() or mark_inode_dirty_sync() instead of calling this directly.

CAREFUL! We only add the inode to the dirty list if it is hashed or if it refers to a blockdev. Unhashed inodes will never be added to the dirty list even if they are later hashed, as they will have been marked dirty already.

In short, ensure you hash any inodes _before_ you start marking them dirty.

Note that for blockdevs, inode->dirtied_when represents the dirtying time of the block-special inode (/dev/hda1) itself. And the ->dirtied_when field of the kernel-internal blockdev inode represents the dirtying time of the blockdev's pages. This is why for I_DIRTY_PAGES we always use page->mapping->host, so the page-dirtying time is recorded in the internal blockdev inode.

void **writeback_inodes_sb_nr**(struct super_block *sb, unsigned long nr, enum wb_reason reason)

writeback dirty inodes from given super_block

**Parameters**

**struct super_block *sb**
the superblock

**unsigned long nr**
the number of pages to write

**enum wb_reason reason**
reason why some writeback work initiated

**Description**

Start writeback on some inodes on this super_block. No guarantees are made on how many (if any) will be written, and this function does not wait for IO completion of submitted IO.

void **writeback_inodes_sb**(struct super_block *sb, enum wb_reason reason)
writeback dirty inodes from given super_block

**Parameters**

**struct super_block *sb**
the superblock

**enum wb_reason reason**
reason why some writeback work was initiated

**Description**

Start writeback on some inodes on this super_block. No guarantees are made on how many (if any) will be written, and this function does not wait for IO completion of submitted IO.

void **try_to_writeback_inodes_sb**(struct super_block *sb, enum wb_reason reason)
try to start writeback if none underway

**Parameters**

**struct super_block *sb**
the superblock

**enum wb_reason reason**
reason why some writeback work was initiated

**Description**

Invoke __writeback_inodes_sb_nr if no writeback is currently underway.

void **sync_inodes_sb**(struct super_block *sb)
sync sb inode pages

**Parameters**

**struct super_block *sb**
the superblock

**Description**

This function writes and waits on any dirty inode belonging to this super_block.

int **write_inode_now**(struct *inode* *inode, int sync)

write an inode to disk

**Parameters**

**struct inode *inode**
inode to write to disk

**int sync**
whether the write should be synchronous or not

**Description**

This function commits an inode to disk immediately if it is dirty. This is primarily needed by knfsd.

The caller must either have a ref on the inode or must have set I_WILL_FREE.

int **sync_inode_metadata**(struct *inode* *inode, int wait)

write an inode to disk

**Parameters**

**struct inode *inode**
the inode to sync

**int wait**
wait for I/O to complete.

**Description**

Write an inode to disk and adjust its dirty state after completion.

**Note**

only writes the actual inode, no associated data or other metadata.

struct file *__anon_inode_getfile__(const char *name, const struct file_operations *fops, void *priv, int flags)

creates a new file instance by hooking it up to an anonymous inode, and a dentry that describe the "class" of the file

**Parameters**

**const char *name**
[in] name of the "class" of the new file

**const struct file_operations *fops**
[in] file operations for the new file

**void *priv**
[in] private data for the new file (will be file's private_data)

**int flags**
[in] flags

**Description**

Creates a new file by hooking it on a single inode. This is useful for files that do not need to have a full-fledged inode in order to operate correctly. All the files created with *anon_inode_getfile()* will share a single inode, hence saving memory and avoiding code duplication for the file/inode/dentry setup. Returns the newly created file* or an error pointer.

struct file *__anon_inode_create_getfile__(const char *name, const struct file_operations *fops, void *priv, int flags, const struct inode *context_inode)

Like *anon_inode_getfile()*, but creates a new !S_PRIVATE anon inode rather than reuse the singleton anon inode and calls the inode_init_security_anon() LSM hook.

**Parameters**

**const char *name**
[in] name of the ”class” of the new file

**const struct file_operations *fops**
[in] file operations for the new file

**void *priv**
[in] private data for the new file (will be file's private_data)

**int flags**
[in] flags

`const struct inode *context_inode`

[in] the logical relationship with the new inode (optional)

**Description**

Create a new anonymous inode and file pair. This can be done for two reasons:

- for the inode to have its own security context, so that LSMs can enforce policy on the inode's creation;
- if the caller needs a unique inode, for example in order to customize the size returned by fstat()

The LSM may use **context_inode** in inode_init_security_anon(), but a reference to it is not held.

Returns the newly created file* or an error pointer.

int __anon_inode_getfd__(const char *name, const struct file_operations *fops, void *priv, int flags)

creates a new file instance by hooking it up to an anonymous inode and a dentry that describe the ”class” of the file

**Parameters**

**const char *name**
[in] name of the ”class” of the new file

**const struct file_operations *fops**
[in] file operations for the new file

**void *priv**
[in] private data for the new file (will be file's private_data)

**int flags**
[in] flags

**Description**

Creates a new file by hooking it on a single inode. This is useful for files that do not need to have a full-fledged inode in order to operate correctly. All the files created with *anon_inode_getfd()*

will use the same singleton inode, reducing memory use and avoiding code duplication for the file/inode/dentry setup. Returns a newly created file descriptor or an error code.

int **setattr_should_drop_sgid**(struct mnt_idmap *idmap, const struct *inode* *inode)
determine whether the setgid bit needs to be removed

**Parameters**

**struct mnt_idmap *idmap**
idmap of the mount **inode** was found from

**const struct inode *inode**
inode to check

**Description**

This function determines whether the setgid bit needs to be removed. We retain backwards compatibility and require setgid bit to be removed unconditionally if S_IXGRP is set. Otherwise we have the exact same requirements as *setattr_prepare()* and *setattr_copy()*.

**Return**

ATTR_KILL_SGID if setgid bit needs to be removed, 0 otherwise.

int **setattr_should_drop_suidgid**(struct mnt_idmap *idmap, struct *inode* *inode)
determine whether the set{g,u}id bit needs to be dropped

**Parameters**

**struct mnt_idmap *idmap**
idmap of the mount **inode** was found from

**struct inode *inode**
inode to check

**Description**

This function determines whether the set{g,u}id bits need to be removed. If the setuid bit needs to be removed ATTR_KILL_SUID is returned. If the setgid bit needs to be removed ATTR_KILL_SGID is returned. If both set{g,u}id bits need to be removed the corresponding mask of both flags is returned.

**Return**

A mask of ATTR_KILL_S{G,U}ID indicating which - if any - setid bits to remove, 0 otherwise.

int **setattr_prepare**(struct mnt_idmap *idmap, struct *dentry* *dentry, struct iattr *attr)
check if attribute changes to a dentry are allowed

**Parameters**

**struct mnt_idmap *idmap**
idmap of the mount the inode was found from

**struct dentry *dentry**
dentry to check

**struct iattr *attr**
attributes to change

### Description

Check if we are allowed to change the attributes contained in **attr** in the given dentry. This includes the normal unix access permission checks, as well as checks for rlimits and others. The function also clears SGID bit from mode if user is not allowed to set it. Also file capabilities and IMA extended attributes are cleared if ATTR_KILL_PRIV is set.

If the inode has been found through an idmapped mount the idmap of the vfsmount must be passed through **idmap**. This function will then take care to map the inode according to **idmap** before checking permissions. On non-idmapped mounts or if permission checking is to be performed on the raw inode simply pass **nop_mnt_idmap**.

Should be called as the first thing in ->setattr implementations, possibly after taking additional locks.

int **inode_newsize_ok**(const struct *inode* *inode, loff_t offset)

may this inode be truncated to a given size

### Parameters

**const struct inode *inode**
the inode to be truncated

**loff_t offset**
the new size to assign to the inode

### Description

inode_newsize_ok must be called with i_mutex held.

inode_newsize_ok will check filesystem limits and ulimits to check that the new inode size is within limits. inode_newsize_ok will also send SIGXFSZ when necessary. Caller must not proceed with inode size change if failure is returned. **inode** must be a file (not directory), with appropriate permissions to allow truncate (inode_newsize_ok does NOT check these conditions).

### Return

0 on success, -ve errno on failure

void **setattr_copy**(struct mnt_idmap *idmap, struct *inode* *inode, const struct iattr *attr)

copy simple metadata updates into the generic inode

### Parameters

**struct mnt_idmap *idmap**
idmap of the mount the inode was found from

**struct inode *inode**
the inode to be updated

**const struct iattr *attr**
the new attributes

### Description

setattr_copy must be called with i_mutex held.

setattr_copy updates the inode's metadata with that specified in attr on idmapped mounts. Necessary permission checks to determine whether or not the S_ISGID property needs to be removed are performed with the correct idmapped mount permission helpers. Noticeably missing is inode size update, which is more complex as it requires pagecache updates.

If the inode has been found through an idmapped mount the idmap of the vfsmount must be passed through **idmap**. This function will then take care to map the inode according to **idmap** before checking permissions. On non-idmapped mounts or if permission checking is to be performed on the raw inode simply pass **nop_mnt_idmap**.

The inode is not marked as dirty after this operation. The rationale is that for "simple" filesystems, the struct inode is the inode storage. The caller is free to mark the inode dirty afterwards if needed.

int **notify_change**(struct mnt_idmap *idmap, struct *dentry* *dentry, struct iattr *attr, struct inode **delegated_inode)

modify attributes of a filesytem object

**Parameters**

**struct mnt_idmap *idmap**
idmap of the mount the inode was found from

**struct dentry *dentry**
object affected

**struct iattr *attr**
new attributes

**struct inode **delegated_inode**
returns inode, if the inode is delegated

**Description**

The caller must hold the i_mutex on the affected object.

If notify_change discovers a delegation in need of breaking, it will return -EWOULDBLOCK and return a reference to the inode in delegated_inode. The caller should then break the delegation and retry. Because breaking a delegation may take a long time, the caller should drop the i_mutex before doing so.

Alternatively, a caller may pass NULL for delegated_inode. This may be appropriate for callers that expect the underlying filesystem not to be NFS exported. Also, passing NULL is fine for callers holding the file open for write, as there can be no conflicting delegation in that case.

If the inode has been found through an idmapped mount the idmap of the vfsmount must be passed through **idmap**. This function will then take care to map the inode according to **idmap** before checking permissions. On non-idmapped mounts or if permission checking is to be performed on the raw inode simply pass **nop_mnt_idmap**.

char ***d_path**(const struct *path* *path, char *buf, int buflen)

return the path of a dentry

**Parameters**

**const struct path *path**
path to report

**char *buf**
buffer to return value in

**int buflen**
buffer length

### Description

Convert a dentry into an ASCII path name. If the entry has been deleted the string ” (deleted)” is appended. Note that this is ambiguous.

Returns a pointer into the buffer or an error code if the path was too long. Note: Callers should use the returned pointer, not the passed in buffer, to use the name! The implementation often starts at an offset into the buffer, and may leave 0 bytes at the start.

”buflen” should be positive.

struct page *__dax_layout_busy_page_range__(struct *address_space* *mapping, loff_t start, loff_t end)

find first pinned page in **mapping**

### Parameters

**struct address_space *mapping**
address space to scan for a page with ref count > 1

**loff_t start**
Starting offset. Page containing 'start' is included.

**loff_t end**
End offset. Page containing 'end' is included. If 'end' is LLONG_MAX, pages from 'start' till the end of file are included.

### Description

DAX requires ZONE_DEVICE mapped pages. These pages are never 'onlined' to the page allocator so they are considered idle when page->count == 1. A filesystem uses this interface to determine if any page in the mapping is busy, i.e. for DMA, or other get_user_pages() usages.

It is expected that the filesystem is holding locks to block the establishment of new mappings in this address_space. I.e. it expects to be able to run unmap_mapping_range() and subsequently not race mapping_mapped() becoming true.

ssize_t __dax_iomap_rw__(struct kiocb *iocb, struct iov_iter *iter, const struct iomap_ops *ops)

Perform I/O to a DAX file

### Parameters

**struct kiocb *iocb**
The control block for this I/O

**struct iov_iter *iter**
The addresses to do I/O from or to

**const struct iomap_ops *ops**
iomap ops passed from the file system

### Description

This function performs read and write operations to directly mapped persistent memory. The callers needs to take care of read/write exclusion and evicting any page cache pages in the region under I/O.

vm_fault_t __dax_iomap_fault__(struct vm_fault *vmf, unsigned int order, pfn_t *pfnp, int *iomap_errp, const struct iomap_ops *ops)

handle a page fault on a DAX file

**Parameters**

**struct vm_fault *vmf**
The description of the fault

**unsigned int order**
Order of the page to fault in

**pfn_t *pfnp**
PFN to insert for synchronous faults if fsync is required

**int *iomap_errp**
Storage for detailed error code in case of error

**const struct iomap_ops *ops**
Iomap ops passed from the file system

**Description**

When a page fault occurs, filesystems may call this helper in their fault handler for DAX files. *dax_iomap_fault()* assumes the caller has done all the necessary locking for page fault to proceed successfully.

vm_fault_t **dax_finish_sync_fault**(struct vm_fault *vmf, unsigned int order, pfn_t pfn)
finish synchronous page fault

**Parameters**

**struct vm_fault *vmf**
The description of the fault

**unsigned int order**
Order of entry to be inserted

**pfn_t pfn**
PFN to insert

**Description**

This function ensures that the file range touched by the page fault is stored persistently on the media and handles inserting of appropriate page table entry.

void **simple_rename_timestamp**(struct inode *old_dir, struct dentry *old_dentry, struct inode *new_dir, struct dentry *new_dentry)
update the various inode timestamps for rename

**Parameters**

**struct inode *old_dir**
old parent directory

**struct dentry *old_dentry**
dentry that is being renamed

**struct inode *new_dir**
new parent directory

**struct dentry *new_dentry**
target for rename

### Description

POSIX mandates that the old and new parent directories have their ctime and mtime updated, and that inodes of **old_dentry** and **new_dentry** (if any), have their ctime updated.

int **simple_setattr**(struct mnt_idmap *idmap, struct *dentry* *dentry, struct *iattr* *iattr)

setattr for simple filesystem

### Parameters

**struct mnt_idmap *idmap**
idmap of the target mount

**struct dentry *dentry**
dentry

**struct iattr *iattr**
iattr structure

### Description

Returns 0 on success, -error on failure.

simple_setattr is a simple ->setattr implementation without a proper implementation of size changes.

It can either be used for in-memory filesystems or special files on simple regular filesystems. Anything that needs to change on-disk or wire state on size changes needs its own setattr method.

ssize_t **simple_read_from_buffer**(void __user *to, size_t count, loff_t *ppos, const void *from, size_t available)

copy data from the buffer to user space

### Parameters

**void __user *to**
the user space buffer to read to

**size_t count**
the maximum number of bytes to read

**loff_t *ppos**
the current position in the buffer

**const void *from**
the buffer to read from

**size_t available**
the size of the buffer

### Description

The *simple_read_from_buffer()* function reads up to **count** bytes from the buffer **from** at offset **ppos** into the user space address starting at **to**.

On success, the number of bytes read is returned and the offset **ppos** is advanced by this number, or negative value is returned on error.

ssize_t **simple_write_to_buffer**(void *to, size_t available, loff_t *ppos, const void __user *from, size_t count)

copy data from user space to the buffer

**Parameters**

**void *to**
the buffer to write to

**size_t available**
the size of the buffer

**loff_t *ppos**
the current position in the buffer

**const void __user *from**
the user space buffer to read from

**size_t count**
the maximum number of bytes to read

**Description**

The *simple_write_to_buffer()* function reads up to **count** bytes from the user space address starting at **from** into the buffer **to** at offset **ppos**.

On success, the number of bytes written is returned and the offset **ppos** is advanced by this number, or negative value is returned on error.

ssize_t **memory_read_from_buffer**(void *to, size_t count, loff_t *ppos, const void *from, size_t available)

copy data from the buffer

**Parameters**

**void *to**
the kernel space buffer to read to

**size_t count**
the maximum number of bytes to read

**loff_t *ppos**
the current position in the buffer

**const void *from**
the buffer to read from

**size_t available**
the size of the buffer

**Description**

The *memory_read_from_buffer()* function reads up to **count** bytes from the buffer **from** at offset **ppos** into the kernel space address starting at **to**.

On success, the number of bytes read is returned and the offset **ppos** is advanced by this number, or negative value is returned on error.

int **generic_encode_ino32_fh**(struct *inode* *inode, __u32 *fh, int *max_len, struct *inode* *parent)

generic export_operations->encode_fh function

**Parameters**

**struct inode *inode**
the object to encode

**__u32 *fh**
where to store the file handle fragment

**int *max_len**
maximum length to store there (in 4 byte units)

**struct inode *parent**
parent directory inode, if wanted

**Description**

This generic encode_fh function assumes that the 32 inode number is suitable for locating an inode, and that the generation number can be used to check that it is still valid. It places them in the filehandle fragment where export_decode_fh expects to find them.

struct dentry *__generic_fh_to_dentry__(struct super_block *sb, struct *fid* *fid, int fh_len, int fh_type, struct inode *(*get_inode)(struct super_block *sb, u64 ino, u32 gen))

generic helper for the fh_to_dentry export operation

**Parameters**

**struct super_block *sb**
filesystem to do the file handle conversion on

**struct fid *fid**
file handle to convert

**int fh_len**
length of the file handle in bytes

**int fh_type**
type of file handle

**struct inode *(*get_inode) (struct super_block *sb, u64 ino, u32 gen)**
filesystem callback to retrieve inode

**Description**

This function decodes **fid** as long as it has one of the well-known Linux filehandle types and calls **get_inode** on it to retrieve the inode for the object specified in the file handle.

struct dentry *__generic_fh_to_parent__(struct super_block *sb, struct *fid* *fid, int fh_len, int fh_type, struct inode *(*get_inode)(struct super_block *sb, u64 ino, u32 gen))

generic helper for the fh_to_parent export operation

**Parameters**

**struct super_block *sb**
filesystem to do the file handle conversion on

**struct fid *fid**
file handle to convert

**int fh_len**
length of the file handle in bytes

**int fh_type**
type of file handle

**struct inode *(*get_inode) (struct super_block *sb, u64 ino, u32 gen)**
filesystem callback to retrieve inode

**Description**

This function decodes **fid** as long as it has one of the well-known Linux filehandle types and calls **get_inode** on it to retrieve the inode for the _parent_ object specified in the file handle if it is specified in the file handle, or NULL otherwise.

int **__generic_file_fsync**(struct *file* *file, loff_t start, loff_t end, int datasync)
generic fsync implementation for simple filesystems

**Parameters**

**struct file *file**
file to synchronize

**loff_t start**
start offset in bytes

**loff_t end**
end offset in bytes (inclusive)

**int datasync**
only synchronize essential metadata if true

**Description**

This is a generic implementation of the fsync method for simple filesystems which track all non-inode metadata in the buffers list hanging off the address_space structure.

int **generic_file_fsync**(struct *file* *file, loff_t start, loff_t end, int datasync)
generic fsync implementation for simple filesystems with flush

**Parameters**

**struct file *file**
file to synchronize

**loff_t start**
start offset in bytes

**loff_t end**
end offset in bytes (inclusive)

**int datasync**
only synchronize essential metadata if true

int **generic_check_addressable**(unsigned blocksize_bits, u64 num_blocks)
Check addressability of file system

**Parameters**

**unsigned blocksize_bits**
log of file system block size

**u64 num_blocks**
number of blocks in file system

**Description**

Determine whether a file system with **num_blocks** blocks (and a block size of 2****blocksize_bits**) is addressable by the sector_t and page cache of the system. Return 0 if so and -EFBIG otherwise.

int **simple_nosetlease**(struct file *filp, int arg, struct file_lock **flp, void **priv)
generic helper for prohibiting leases

**Parameters**

**struct file *filp**
file pointer

**int arg**
type of lease to obtain

**struct file_lock **flp**
new lease supplied for insertion

**void **priv**
private data for lm_setup operation

**Description**

Generic helper for filesystems that do not wish to allow leases to be set. All arguments are ignored and it just returns -EINVAL.

const char ***simple_get_link**(struct *dentry* *dentry, struct *inode* *inode, struct delayed_call *done)
generic helper to get the target of ”fast” symlinks

**Parameters**

**struct dentry *dentry**
not used here

**struct inode *inode**
the symlink inode

**struct delayed_call *done**
not used here

**Description**

Generic helper for filesystems to use for symlink inodes where a pointer to the symlink target is stored in ->i_link. NOTE: this isn't normally called, since as an optimization the path lookup code uses any non-NULL ->i_link directly, without calling ->get_link(). But ->get_link() still must be set, to mark the inode_operations as being for a symlink.

**Return**

the symlink target

void **generic_set_encrypted_ci_d_ops**(struct *dentry* *dentry)

helper for setting d_ops for given dentry

**Parameters**

**struct dentry *dentry**
dentry to set ops on

**Description**

Casefolded directories need d_hash and d_compare set, so that the dentries contained in them are handled case-insensitively. Note that these operations are needed on the parent directory rather than on the dentries in it, and while the casefolding flag can be toggled on and off on an empty directory, dentry_operations can't be changed later. As a result, if the filesystem has casefolding support enabled at all, we have to give all dentries the casefolding operations even if their inode doesn't have the casefolding flag currently (and thus the casefolding ops would be no-ops for now).

Encryption works differently in that the only dentry operation it needs is d_revalidate, which it only needs on dentries that have the no-key name flag. The no-key flag can't be set ”later”, so we don't have to worry about that.

Finally, to maximize compatibility with overlayfs (which isn't compatible with certain dentry operations) and to avoid taking an unnecessary performance hit, we use custom dentry_operations for each possible combination rather than always installing all operations.

bool **inode_maybe_inc_iversion**(struct *inode* *inode, bool force)

increments i_version

**Parameters**

**struct inode *inode**
inode with the i_version that should be updated

**bool force**
increment the counter even if it's not necessary?

**Description**

Every time the inode is modified, the i_version field must be seen to have changed by any observer.

If ”force” is set or the QUERIED flag is set, then ensure that we increment the value, and clear the queried flag.

In the common case where neither is set, then we can return ”false” without updating i_version.

If this function returns false, and no other metadata has changed, then we can avoid logging the metadata.

u64 **inode_query_iversion**(struct *inode* *inode)

read i_version for later use

**Parameters**

**struct inode *inode**
inode from which i_version should be read

**Description**

Read the inode i_version counter. This should be used by callers that wish to store the returned i_version for later comparison. This will guarantee that a later query of the i_version will result in a different value if anything has changed.

In this implementation, we fetch the current value, set the QUERIED flag and then try to swap it into place with a cmpxchg, if it wasn't already set. If that fails, we try again with the newly fetched value from the cmpxchg.

struct timespec64 **simple_inode_init_ts**(struct *inode* *inode)

initialize the timestamps for a new inode

**Parameters**

**struct inode *inode**
inode to be initialized

**Description**

When a new inode is created, most filesystems set the timestamps to the current time. Add a helper to do this.

int **posix_acl_chmod**(struct mnt_idmap *idmap, struct *dentry* *dentry, umode_t mode)

chmod a posix acl

**Parameters**

**struct mnt_idmap *idmap**
idmap of the mount **inode** was found from

**struct dentry *dentry**
dentry to check permissions on

**umode_t mode**
the new mode of **inode**

**Description**

If the dentry has been found through an idmapped mount the idmap of the vfsmount must be passed through **idmap**. This function will then take care to map the inode according to **idmap** before checking permissions. On non-idmapped mounts or if permission checking is to be performed on the raw inode simply pass **nop_mnt_idmap**.

int **posix_acl_update_mode**(struct mnt_idmap *idmap, struct *inode* *inode, umode_t *mode_p, struct posix_acl **acl)

update mode in set_acl

**Parameters**

**struct mnt_idmap *idmap**
idmap of the mount **inode** was found from

**struct inode *inode**
target inode

**umode_t *mode_p**
mode (pointer) for update

**struct posix_acl **acl**
acl pointer

**Description**

Update the file mode when setting an ACL: compute the new file permission bits based on the ACL. In addition, if the ACL is equivalent to the new file mode, set ***acl** to NULL to indicate that no ACL should be set.

As with chmod, clear the setgid bit if the caller is not in the owning group or capable of CAP_FSETID (see inode_change_ok).

If the inode has been found through an idmapped mount the idmap of the vfsmount must be passed through **idmap**. This function will then take care to map the inode according to **idmap** before checking permissions. On non-idmapped mounts or if permission checking is to be performed on the raw inode simply pass **nop_mnt_idmap**.

Called from set_acl inode operations.

struct posix_acl *__posix_acl_from_xattr__(struct user_namespace *userns, const void *value, size_t size)

convert POSIX ACLs from backing store to VFS format

**Parameters**

**struct user_namespace *userns**
the filesystem's idmapping

**const void *value**
the uapi representation of POSIX ACLs

**size_t size**
the size of **void**

**Description**

Filesystems that store POSIX ACLs in the unaltered uapi format should use *posix_acl_from_xattr()* when reading them from the backing store and converting them into the struct posix_acl VFS format. The helper is specifically intended to be called from the acl inode operation.

The *posix_acl_from_xattr()* function will map the raw {g,u}id values stored in ACL_{GROUP,USER} entries into idmapping in **userns**.

Note that *posix_acl_from_xattr()* does not take idmapped mounts into account. If it did it calling it from the get acl inode operation would return POSIX ACLs mapped according to an idmapped mount which would mean that the value couldn't be cached for the filesystem. Idmapped mounts are taken into account on the fly during permission checking or right at the VFS - userspace boundary before reporting them to the user.

**Return**

**Allocated struct posix_acl on success, NULL for a valid header but**
without actual POSIX ACL entries, or ERR_PTR() encoded error code.

int __vfs_set_acl__(struct mnt_idmap *idmap, struct *dentry* *dentry, const char *acl_name, struct posix_acl *kacl)

set posix acls

**Parameters**

**struct mnt_idmap *idmap**
idmap of the mount

**struct dentry *dentry**
: the dentry based on which to set the posix acls

**const char *acl_name**
: the name of the posix acl

**struct posix_acl *kacl**
: the posix acls in the appropriate VFS format

**Description**

This function sets **kacl**. The caller must all posix_acl_release() on **kacl** afterwards.

**Return**

On success 0, on error negative errno.

struct posix_acl ***vfs_get_acl**(struct mnt_idmap *idmap, struct *dentry* *dentry, const char *acl_name)

: get posix acls

**Parameters**

**struct mnt_idmap *idmap**
: idmap of the mount

**struct dentry *dentry**
: the dentry based on which to retrieve the posix acls

**const char *acl_name**
: the name of the posix acl

**Description**

This function retrieves **kacl** from the filesystem. The caller must all posix_acl_release() on **kacl**.

**Return**

On success POSIX ACLs in VFS format, on error negative errno.

int **vfs_remove_acl**(struct mnt_idmap *idmap, struct *dentry* *dentry, const char *acl_name)

: remove posix acls

**Parameters**

**struct mnt_idmap *idmap**
: idmap of the mount

**struct dentry *dentry**
: the dentry based on which to retrieve the posix acls

**const char *acl_name**
: the name of the posix acl

**Description**

This function removes posix acls.

**Return**

On success 0, on error negative errno.

void **generic_fillattr**(struct mnt_idmap *idmap, u32 request_mask, struct *inode* *inode, struct kstat *stat)

Fill in the basic attributes from the inode struct

**Parameters**

**struct mnt_idmap *idmap**
idmap of the mount the inode was found from

**u32 request_mask**
statx request_mask

**struct inode *inode**
Inode to use as the source

**struct kstat *stat**
Where to fill in the attributes

**Description**

Fill in the basic attributes in the kstat structure from data that's to be found on the VFS inode structure. This is the default if no getattr inode operation is supplied.

If the inode has been found through an idmapped mount the idmap of the vfsmount must be passed through **idmap**. This function will then take care to map the inode according to **idmap** before filling in the uid and gid filds. On non-idmapped mounts or if permission checking is to be performed on the raw inode simply pass **nop_mnt_idmap**.

void **generic_fill_statx_attr**(struct *inode* *inode, struct kstat *stat)

Fill in the statx attributes from the inode flags

**Parameters**

**struct inode *inode**
Inode to use as the source

**struct kstat *stat**
Where to fill in the attribute flags

**Description**

Fill in the STATX_ATTR_* flags in the kstat structure for properties of the inode that are published on i_flags and enforced by the VFS.

int **vfs_getattr_nosec**(const struct *path* *path, struct kstat *stat, u32 request_mask, unsigned int query_flags)

getattr without security checks

**Parameters**

**const struct path *path**
file to get attributes from

**struct kstat *stat**
structure to return attributes in

**u32 request_mask**
STATX_xxx flags indicating what the caller wants

**unsigned int query_flags**
Query mode (AT_STATX_SYNC_TYPE)

**Description**

Get attributes without calling security_inode_getattr.

Currently the only caller other than vfs_getattr is internal to the filehandle lookup code, which uses only the inode number and returns no attributes to any user. Any other code probably wants vfs_getattr.

int **vfs_fsync_range**(struct *file* *file, loff_t start, loff_t end, int datasync)

helper to sync a range of data & metadata to disk

**Parameters**

**struct file *file**
file to sync

**loff_t start**
offset in bytes of the beginning of data range to sync

**loff_t end**
offset in bytes of the end of data range (inclusive)

**int datasync**
perform only datasync

**Description**

Write back data in range **start**..**end** and metadata for **file** to disk. If **datasync** is set only metadata needed to access modified file data is written.

int **vfs_fsync**(struct *file* *file, int datasync)

perform a fsync or fdatasync on a file

**Parameters**

**struct file *file**
file to sync

**int datasync**
only perform a fdatasync operation

**Description**

Write back data and metadata for **file** to disk. If **datasync** is set only metadata needed to access modified file data is written.

int **__vfs_setxattr_locked**(struct mnt_idmap *idmap, struct *dentry* *dentry, const char *name, const void *value, size_t size, int flags, struct inode **delegated_inode)

set an extended attribute while holding the inode lock

**Parameters**

**struct mnt_idmap *idmap**
idmap of the mount of the target inode

**struct dentry *dentry**
object to perform setxattr on

**const char *name**
xattr name to set

**const void *value**
value to set **name** to

**size_t size**
size of **value**

**int flags**
flags to pass into filesystem operations

**struct inode **delegated_inode**
on return, will contain an inode pointer that a delegation was broken on, NULL if none.

ssize_t **vfs_listxattr**(struct *dentry* *dentry, char *list, size_t size)
retrieve 0 separated list of xattr names

**Parameters**

**struct dentry *dentry**
the dentry from whose inode the xattr names are retrieved

**char *list**
buffer to store xattr names into

**size_t size**
size of the buffer

**Description**

This function returns the names of all xattrs associated with the inode of **dentry**.

Note, for legacy reasons the *vfs_listxattr()* function lists POSIX ACLs as well. Since POSIX ACLs are decoupled from IOP_XATTR the *vfs_listxattr()* function doesn't check for this flag since a filesystem could implement POSIX ACLs without implementing any other xattrs.

However, since all codepaths that remove IOP_XATTR also assign of inode operations that either don't implement or implement a stub ->listxattr() operation.

**Return**

**On success, the size of the buffer that was used. On error a**
negative error code.

int **__vfs_removexattr_locked**(struct mnt_idmap *idmap, struct *dentry* *dentry, const char *name, struct inode **delegated_inode)
set an extended attribute while holding the inode lock

**Parameters**

**struct mnt_idmap *idmap**
idmap of the mount of the target inode

**struct dentry *dentry**
object to perform setxattr on

**const char *name**
name of xattr to remove

**struct inode **delegated_inode**
on return, will contain an inode pointer that a delegation was broken on, NULL if none.

ssize_t **generic_listxattr**(struct *dentry* *dentry, char *buffer, size_t buffer_size)
run through a dentry's xattr list() operations

**Parameters**

**struct dentry *dentry**
dentry to list the xattrs

**char *buffer**
result buffer

**size_t buffer_size**
size of **buffer**

**Description**

Combine the results of the list() operation from every xattr_handler in the xattr_handler stack.

Note that this will not include the entries for POSIX ACLs.

const char ***xattr_full_name**(const struct xattr_handler *handler, const char *name)
Compute full attribute name from suffix

**Parameters**

**const struct xattr_handler *handler**
handler of the xattr_handler operation

**const char *name**
name passed to the xattr_handler operation

**Description**

The get and set xattr handler operations are called with the remainder of the attribute name after skipping the handler's prefix: for example, ”foo” is passed to the get operation of a handler with prefix ”user.” to get attribute ”user.foo”. The full name is still ”there” in the name though.

**Note**

the list xattr handler operation when called from the vfs is passed a NULL name; some file systems use this operation internally, with varying semantics.

int **mnt_get_write_access**(struct vfsmount *m)
get write access to a mount without freeze protection

**Parameters**

**struct vfsmount *m**
the mount on which to take a write

**Description**

This tells the low-level filesystem that a write is about to be performed to it, and makes sure that writes are allowed (mnt it read-write) before returning success. This operation does not protect against filesystem being frozen. When the write operation is finished, *mnt_put_write_access()* must be called. This is effectively a refcount.

int **mnt_want_write**(struct vfsmount *m)
get write access to a mount

**Parameters**

**struct vfsmount *m**
the mount on which to take a write

**Description**

This tells the low-level filesystem that a write is about to be performed to it, and makes sure that writes are allowed (mount is read-write, filesystem is not frozen) before returning success. When the write operation is finished, *mnt_drop_write()* must be called. This is effectively a refcount.

int **mnt_want_write_file**(struct *file* *file)
get write access to a file's mount

**Parameters**

**struct file *file**
the file who's mount on which to take a write

**Description**

This is like mnt_want_write, but if the file is already open for writing it skips incrementing mnt_writers (since the open file already has a reference) and instead only does the freeze protection and the check for emergency r/o remounts. This must be paired with mnt_drop_write_file.

void **mnt_put_write_access**(struct vfsmount *mnt)
give up write access to a mount

**Parameters**

**struct vfsmount *mnt**
the mount on which to give up write access

**Description**

Tells the low-level filesystem that we are done performing writes to it. Must be matched with *mnt_get_write_access()* call above.

void **mnt_drop_write**(struct vfsmount *mnt)
give up write access to a mount

**Parameters**

**struct vfsmount *mnt**
the mount on which to give up write access

**Description**

Tells the low-level filesystem that we are done performing writes to it and also allows filesystem to be frozen again. Must be matched with *mnt_want_write()* call above.

struct vfsmount ***vfs_create_mount**(struct fs_context *fc)
Create a mount for a configured superblock

**Parameters**

**struct fs_context *fc**
The configuration context with the superblock attached

**Description**

Create a mount to an already configured superblock. If necessary, the caller should invoke *vfs_get_tree()* before calling this.

Note that this does not attach the mount to anything.

bool **path_is_mountpoint**(const struct *path* *path)

Check if path is a mount in the current namespace.

**Parameters**

**const struct path *path**
path to check

d_mountpoint() can only be used reliably to establish if a dentry is not mounted in any namespace and that common case is handled inline. d_mountpoint() isn't aware of the possibility there may be multiple mounts using a given dentry in a different namespace. This function checks if the passed in path is a mountpoint rather than the dentry alone.

int **may_umount_tree**(struct vfsmount *m)

check if a mount tree is busy

**Parameters**

**struct vfsmount *m**
root of mount tree

**Description**

This is called to check if a tree of mounts has any open files, pwds, chroots or sub mounts that are busy.

int **may_umount**(struct vfsmount *mnt)

check if a mount point is busy

**Parameters**

**struct vfsmount *mnt**
root of mount

**Description**

This is called to check if a mount point has any open files, pwds, chroots or sub mounts. If the mount has sub mounts this will return busy regardless of whether the sub mounts are busy.

Doesn't take quota and stuff into account. IOW, in some cases it will give false negatives. The main reason why it's here is that we need a non-destructive way to look for easily umountable filesystems.

struct vfsmount ***clone_private_mount**(const struct *path* *path)

create a private clone of a path

**Parameters**

**const struct path *path**
path to clone

**Description**

This creates a new vfsmount, which will be the clone of **path**. The new mount will not be attached anywhere in the namespace and will be private (i.e. changes to the originating mount won't be propagated into this).

Release with mntput().

void **mnt_set_expiry**(struct vfsmount *mnt, struct list_head *expiry_list)
Put a mount on an expiration list

**Parameters**

**struct vfsmount *mnt**
The mount to list.

**struct list_head *expiry_list**
The list to add the mount to.

### 1.3.2 The proc filesystem

#### sysctl interface

int **proc_dostring**(struct ctl_table *table, int write, void *buffer, size_t *lenp, loff_t *ppos)
read a string sysctl

**Parameters**

**struct ctl_table *table**
the sysctl table

**int write**
TRUE if this is a write to the sysctl file

**void *buffer**
the user buffer

**size_t *lenp**
the size of the user buffer

**loff_t *ppos**
file position

**Description**

Reads/writes a string from/to the user buffer. If the kernel buffer provided is not large enough to hold the string, the string is truncated. The copied string is `NULL-terminated`. If the string is being read by the user process, it is copied and a newline 'n' is added. It is truncated if the buffer is not large enough.

Returns 0 on success.

int **proc_dobool**(struct ctl_table *table, int write, void *buffer, size_t *lenp, loff_t *ppos)
read/write a bool

**Parameters**

**struct ctl_table *table**
the sysctl table

**int write**
TRUE if this is a write to the sysctl file

**void *buffer**
the user buffer

**size_t *lenp**
the size of the user buffer

**loff_t *ppos**
file position

**Description**

Reads/writes one integer value from/to the user buffer, treated as an ASCII string.

table->data must point to a bool variable and table->maxlen must be sizeof(bool).

Returns 0 on success.

int **proc_dointvec**(struct ctl_table *table, int write, void *buffer, size_t *lenp, loff_t *ppos)
read a vector of integers

**Parameters**

**struct ctl_table *table**
the sysctl table

**int write**
TRUE if this is a write to the sysctl file

**void *buffer**
the user buffer

**size_t *lenp**
the size of the user buffer

**loff_t *ppos**
file position

**Description**

Reads/writes up to table->maxlen/sizeof(unsigned int) integer values from/to the user buffer, treated as an ASCII string.

Returns 0 on success.

int **proc_douintvec**(struct ctl_table *table, int write, void *buffer, size_t *lenp, loff_t *ppos)
read a vector of unsigned integers

**Parameters**

**struct ctl_table *table**
the sysctl table

**int write**
TRUE if this is a write to the sysctl file

**void *buffer**
the user buffer

**size_t *lenp**
the size of the user buffer

**loff_t *ppos**
file position

**Description**

Reads/writes up to table->maxlen/sizeof(unsigned int) unsigned integer values from/to the user buffer, treated as an ASCII string.

Returns 0 on success.

int **proc_dointvec_minmax**(struct ctl_table *table, int write, void *buffer, size_t *lenp, loff_t *ppos)

read a vector of integers with min/max values

**Parameters**

**struct ctl_table *table**
the sysctl table

**int write**
TRUE if this is a write to the sysctl file

**void *buffer**
the user buffer

**size_t *lenp**
the size of the user buffer

**loff_t *ppos**
file position

**Description**

Reads/writes up to table->maxlen/sizeof(unsigned int) integer values from/to the user buffer, treated as an ASCII string.

This routine will ensure the values are within the range specified by table->extra1 (min) and table->extra2 (max).

Returns 0 on success or -EINVAL on write when the range check fails.

int **proc_douintvec_minmax**(struct ctl_table *table, int write, void *buffer, size_t *lenp, loff_t *ppos)

read a vector of unsigned ints with min/max values

**Parameters**

**struct ctl_table *table**
the sysctl table

**int write**
TRUE if this is a write to the sysctl file

**void *buffer**
the user buffer

**size_t *lenp**
the size of the user buffer

**loff_t *ppos**
file position

**Description**

Reads/writes up to table->maxlen/sizeof(unsigned int) unsigned integer values from/to the user buffer, treated as an ASCII string. Negative strings are not allowed.

This routine will ensure the values are within the range specified by table->extra1 (min) and table->extra2 (max). There is a final sanity check for UINT_MAX to avoid having to support wrap around uses from userspace.

Returns 0 on success or -ERANGE on write when the range check fails.

int **proc_dou8vec_minmax**(struct ctl_table *table, int write, void *buffer, size_t *lenp, loff_t *ppos)

read a vector of unsigned chars with min/max values

**Parameters**

**struct ctl_table *table**
the sysctl table

**int write**
TRUE if this is a write to the sysctl file

**void *buffer**
the user buffer

**size_t *lenp**
the size of the user buffer

**loff_t *ppos**
file position

**Description**

Reads/writes up to table->maxlen/sizeof(u8) unsigned chars values from/to the user buffer, treated as an ASCII string. Negative strings are not allowed.

This routine will ensure the values are within the range specified by table->extra1 (min) and table->extra2 (max).

Returns 0 on success or an error on write when the range check fails.

int **proc_doulongvec_minmax**(struct ctl_table *table, int write, void *buffer, size_t *lenp, loff_t *ppos)

read a vector of long integers with min/max values

**Parameters**

**struct ctl_table *table**
the sysctl table

**int write**
TRUE if this is a write to the sysctl file

**void *buffer**
the user buffer

**size_t *lenp**
the size of the user buffer

**loff_t *ppos**
file position

**Description**

Reads/writes up to table->maxlen/sizeof(unsigned long) unsigned long values from/to the user buffer, treated as an ASCII string.

This routine will ensure the values are within the range specified by table->extra1 (min) and table->extra2 (max).

Returns 0 on success.

int **proc_doulongvec_ms_jiffies_minmax**(struct ctl_table *table, int write, void *buffer, size_t *lenp, loff_t *ppos)

read a vector of millisecond values with min/max values

**Parameters**

**struct ctl_table *table**
the sysctl table

**int write**
TRUE if this is a write to the sysctl file

**void *buffer**
the user buffer

**size_t *lenp**
the size of the user buffer

**loff_t *ppos**
file position

**Description**

Reads/writes up to table->maxlen/sizeof(unsigned long) unsigned long values from/to the user buffer, treated as an ASCII string. The values are treated as milliseconds, and converted to jiffies when they are stored.

This routine will ensure the values are within the range specified by table->extra1 (min) and table->extra2 (max).

Returns 0 on success.

int **proc_dointvec_jiffies**(struct ctl_table *table, int write, void *buffer, size_t *lenp, loff_t *ppos)

read a vector of integers as seconds

**Parameters**

**struct ctl_table *table**
the sysctl table

**int write**
TRUE if this is a write to the sysctl file

**void *buffer**
the user buffer

**size_t *lenp**
the size of the user buffer

**loff_t *ppos**
file position

**Description**

Reads/writes up to table->maxlen/sizeof(unsigned int) integer values from/to the user buffer, treated as an ASCII string. The values read are assumed to be in seconds, and are converted into jiffies.

Returns 0 on success.

int **proc_dointvec_userhz_jiffies**(struct ctl_table *table, int write, void *buffer, size_t *lenp, loff_t *ppos)

read a vector of integers as 1/USER_HZ seconds

**Parameters**

**struct ctl_table *table**
the sysctl table

**int write**
TRUE if this is a write to the sysctl file

**void *buffer**
the user buffer

**size_t *lenp**
the size of the user buffer

**loff_t *ppos**
pointer to the file position

**Description**

Reads/writes up to table->maxlen/sizeof(unsigned int) integer values from/to the user buffer, treated as an ASCII string. The values read are assumed to be in 1/USER_HZ seconds, and are converted into jiffies.

Returns 0 on success.

int **proc_dointvec_ms_jiffies**(struct ctl_table *table, int write, void *buffer, size_t *lenp, loff_t *ppos)

read a vector of integers as 1 milliseconds

**Parameters**

**struct ctl_table *table**
the sysctl table

**int write**
TRUE if this is a write to the sysctl file

**void *buffer**
the user buffer

**size_t *lenp**
the size of the user buffer

**loff_t *ppos**
the current position in the file

**Description**

Reads/writes up to table->maxlen/sizeof(unsigned int) integer values from/to the user buffer, treated as an ASCII string. The values read are assumed to be in 1/1000 seconds, and are converted into jiffies.

Returns 0 on success.

int **proc_do_large_bitmap**(struct ctl_table *table, int write, void *buffer, size_t *lenp, loff_t *ppos)

read/write from/to a large bitmap

**Parameters**

**struct ctl_table *table**
the sysctl table

**int write**
TRUE if this is a write to the sysctl file

**void *buffer**
the user buffer

**size_t *lenp**
the size of the user buffer

**loff_t *ppos**
file position

**Description**

The bitmap is stored at table->data and the bitmap length (in bits) in table->maxlen.

We use a range comma separated format (e.g. 1,3-4,10-10) so that large bitmaps may be represented in a compact manner. Writing into the file will clear the bitmap then update it with the given input.

Returns 0 on success.

### proc filesystem interface

void **proc_flush_pid**(struct *pid* *pid)

Remove dcache entries for **pid** from the /proc dcache.

**Parameters**

**struct pid *pid**
pid that should be flushed.

**Description**

This function walks a list of inodes (that belong to any proc filesystem) that are attached to the pid and flushes them from the dentry cache.

It is safe and reasonable to cache /proc entries for a task until that task exits. After that they just clog up the dcache with useless entries, possibly causing useful dcache entries to be flushed instead. This routine is provided to flush those useless dcache entries when a process is reaped.

**NOTE**

**This routine is just an optimization so it does not guarantee**
that no dcache entries will exist after a process is reaped it just makes it very unlikely that any will persist.

### 1.3.3 Events based on file descriptors

void **eventfd_signal_mask**(struct eventfd_ctx *ctx, __poll_t mask)
Increment the event counter

**Parameters**

**struct eventfd_ctx *ctx**
[in] Pointer to the eventfd context.

**__poll_t mask**
[in] poll mask

**Description**

This function is supposed to be called by the kernel in paths that do not allow sleeping. In this function we allow the counter to reach the ULLONG_MAX value, and we signal this as overflow condition by returning a EPOLLERR to poll(2).

void **eventfd_ctx_put**(struct eventfd_ctx *ctx)
Releases a reference to the internal eventfd context.

**Parameters**

**struct eventfd_ctx *ctx**
[in] Pointer to eventfd context.

**Description**

The eventfd context reference must have been previously acquired either with *eventfd_ctx_fdget()* or *eventfd_ctx_fileget()*.

int **eventfd_ctx_remove_wait_queue**(struct eventfd_ctx *ctx, wait_queue_entry_t *wait, __u64 *cnt)
Read the current counter and removes wait queue.

**Parameters**

**struct eventfd_ctx *ctx**
[in] Pointer to eventfd context.

**wait_queue_entry_t *wait**
[in] Wait queue to be removed.

**__u64 *cnt**
[out] Pointer to the 64-bit counter value.

**Description**

Returns `0` if successful, or the following error codes:

**-EAGAIN** : The operation would have blocked.

This is used to atomically remove a wait queue entry from the eventfd wait queue head, and read/reset the counter value.

struct file *__eventfd_fget__(int fd)

Acquire a reference of an eventfd file descriptor.

**Parameters**

**int fd**
[in] Eventfd file descriptor.

**Description**

Returns a pointer to the eventfd file structure in case of success, or the following error pointer:

**-EBADF** : Invalid **fd** file descriptor.

**-EINVAL** : The **fd** file descriptor is not an eventfd file.

struct eventfd_ctx *__eventfd_ctx_fdget__(int fd)

Acquires a reference to the internal eventfd context.

**Parameters**

**int fd**
[in] Eventfd file descriptor.

**Description**

Returns a pointer to the internal eventfd context, otherwise the error pointers returned by the following functions:

eventfd_fget

struct eventfd_ctx *__eventfd_ctx_fileget__(struct *file* *file)

Acquires a reference to the internal eventfd context.

**Parameters**

**struct file *file**
[in] Eventfd file pointer.

**Description**

Returns a pointer to the internal eventfd context, otherwise the error pointer:

**-EINVAL** : The **fd** file descriptor is not an eventfd file.

### 1.3.4 eventpoll (epoll) interfaces

int __ep_events_available__(struct eventpoll *ep)

Checks if ready events might be available.

**Parameters**

**struct eventpoll *ep**
Pointer to the eventpoll context.

**Return**

**a value different than `zero` if ready events are available,**
or `zero` otherwise.

int **reverse_path_check**(void)

The tfile_check_list is list of epitem_head, which have links that are proposed to be newly added. We need to make sure that those added links don't add too many paths such that we will spend all our time waking up eventpoll objects.

**Parameters**

**void**
no arguments

**Return**

**zero if the proposed links don't create too many paths,**
-1 otherwise.

int **ep_poll**(struct eventpoll *ep, struct epoll_event __user *events, int maxevents, struct timespec64 *timeout)

Retrieves ready events, and delivers them to the caller-supplied event buffer.

**Parameters**

**struct eventpoll *ep**
Pointer to the eventpoll context.

**struct epoll_event __user *events**
Pointer to the userspace buffer where the ready events should be stored.

**int maxevents**
Size (in terms of number of events) of the caller event buffer.

**struct timespec64 *timeout**
Maximum timeout for the ready events fetch operation, in timespec. If the timeout is zero, the function will not block, while if the **timeout** ptr is NULL, the function will block until at least one event has been retrieved (or an error occurred).

**Return**

**the number of ready events which have been fetched, or an**
error code, in case of error.

int **ep_loop_check_proc**(struct eventpoll *ep, int depth)

verify that adding an epoll file inside another epoll structure does not violate the constraints, in terms of closed loops, or too deep chains (which can result in excessive stack usage).

**Parameters**

**struct eventpoll *ep**
the `struct eventpoll` to be currently checked.

**int depth**
Current depth of the path being checked.

**Return**

**zero if adding the epoll file inside current epoll**
structure **ep** does not violate the constraints, or -1 otherwise.

int **ep_loop_check**(struct eventpoll *ep, struct eventpoll *to)

Performs a check to verify that adding an epoll file (**to**) into another epoll file (represented by **ep**) does not create closed loops or too deep chains.

**Parameters**

**struct eventpoll *ep**
Pointer to the epoll we are inserting into.

**struct eventpoll *to**
Pointer to the epoll to be inserted.

**Return**

`zero` if adding the epoll **to** inside the epoll **from** does not violate the constraints, or `-1` otherwise.

### 1.3.5 The Filesystem for Exporting Kernel Objects

int **sysfs_create_file_ns**(struct kobject *kobj, const struct attribute *attr, const void *ns)

create an attribute file for an object with custom ns

**Parameters**

**struct kobject *kobj**
object we're creating for

**const struct attribute *attr**
attribute descriptor

**const void *ns**
namespace the new file should belong to

int **sysfs_add_file_to_group**(struct kobject *kobj, const struct attribute *attr, const char *group)

add an attribute file to a pre-existing group.

**Parameters**

**struct kobject *kobj**
object we're acting for.

**const struct attribute *attr**
attribute descriptor.

**const char *group**
group name.

int **sysfs_chmod_file**(struct kobject *kobj, const struct attribute *attr, umode_t mode)

update the modified mode value on an object attribute.

**Parameters**

**struct kobject *kobj**
object we're acting for.

**const struct attribute *attr**
attribute descriptor.

**umode_t mode**
file permissions.

struct kernfs_node *__sysfs_break_active_protection__(struct kobject *kobj, const struct attribute *attr)

break ”active” protection

**Parameters**

**struct kobject *kobj**
The kernel object **attr** is associated with.

**const struct attribute *attr**
The attribute to break the ”active” protection for.

**Description**

With sysfs, just like kernfs, deletion of an attribute is postponed until all active .show() and .store() callbacks have finished unless this function is called. Hence this function is useful in methods that implement self deletion.

void __sysfs_unbreak_active_protection__(struct kernfs_node *kn)

restore ”active” protection

**Parameters**

**struct kernfs_node *kn**
Pointer returned by *sysfs_break_active_protection()*.

**Description**

Undo the effects of *sysfs_break_active_protection()*. Since this function calls kernfs_put() on the kernfs node that corresponds to the 'attr' argument passed to *sysfs_break_active_protection()* that attribute may have been removed between the *sysfs_break_active_protection()* and *sysfs_unbreak_active_protection()* calls, it is not safe to access **kn** after this function has returned.

void __sysfs_remove_file_ns__(struct kobject *kobj, const struct attribute *attr, const void *ns)

remove an object attribute with a custom ns tag

**Parameters**

**struct kobject *kobj**
object we're acting for

**const struct attribute *attr**
attribute descriptor

**const void *ns**
namespace tag of the file to remove

**Description**

Hash the attribute name and namespace tag and kill the victim.

bool __sysfs_remove_file_self__(struct kobject *kobj, const struct attribute *attr)

remove an object attribute from its own method

**Parameters**

**struct kobject *kobj**
object we're acting for

**const struct attribute *attr**
attribute descriptor

**Description**

See kernfs_remove_self() for details.

void **sysfs_remove_file_from_group**(struct kobject *kobj, const struct attribute *attr, const char *group)

remove an attribute file from a group.

**Parameters**

**struct kobject *kobj**
object we're acting for.

**const struct attribute *attr**
attribute descriptor.

**const char *group**
group name.

int **sysfs_create_bin_file**(struct kobject *kobj, const struct bin_attribute *attr)

create binary file for object.

**Parameters**

**struct kobject *kobj**
object.

**const struct bin_attribute *attr**
attribute descriptor.

void **sysfs_remove_bin_file**(struct kobject *kobj, const struct bin_attribute *attr)

remove binary file for object.

**Parameters**

**struct kobject *kobj**
object.

**const struct bin_attribute *attr**
attribute descriptor.

int **sysfs_file_change_owner**(struct kobject *kobj, const char *name, kuid_t kuid, kgid_t kgid)

change owner of a sysfs file.

**Parameters**

**struct kobject *kobj**
object.

**const char *name**
name of the file to change.

**kuid_t kuid**
new owner's kuid

**kgid_t kgid**
new owner's kgid

**Description**

This function looks up the sysfs entry **name** under **kobj** and changes the ownership to **kuid/kgid**.

Returns 0 on success or error code on failure.

int **sysfs_change_owner**(struct kobject *kobj, kuid_t kuid, kgid_t kgid)
change owner of the given object.

**Parameters**

**struct kobject *kobj**
object.

**kuid_t kuid**
new owner's kuid

**kgid_t kgid**
new owner's kgid

**Description**

Change the owner of the default directory, files, groups, and attributes of **kobj** to **kuid/kgid**. Note that sysfs_change_owner mirrors how the sysfs entries for a kobject are added by driver core. In summary, *sysfs_change_owner()* takes care of the default directory entry for **kobj**, the default attributes associated with the ktype of **kobj** and the default attributes associated with the ktype of **kobj**. Additional properties not added by driver core have to be changed by the driver or subsystem which created them. This is similar to how driver/subsystem specific entries are removed.

Returns 0 on success or error code on failure.

int **sysfs_emit**(char *buf, const char *fmt, ...)
scnprintf equivalent, aware of PAGE_SIZE buffer.

**Parameters**

**char *buf**
start of PAGE_SIZE buffer.

**const char *fmt**
format

**...**
optional arguments to **format**

**Description**

Returns number of characters written to **buf**.

int **sysfs_emit_at**(char *buf, int at, const char *fmt, ...)
scnprintf equivalent, aware of PAGE_SIZE buffer.

**Parameters**

**char *buf**
start of PAGE_SIZE buffer.

**int at**
offset in **buf** to start write in bytes **at** must be >= 0 && < PAGE_SIZE

**const char *fmt**
format

**...**
optional arguments to **fmt**

**Description**

Returns number of characters written starting at &**buf**[**at**].

int **sysfs_create_link**(struct kobject *kobj, struct kobject *target, const char *name)
create symlink between two objects.

**Parameters**

**struct kobject *kobj**
object whose directory we're creating the link in.

**struct kobject *target**
object we're pointing to.

**const char *name**
name of the symlink.

int **sysfs_create_link_nowarn**(struct kobject *kobj, struct kobject *target, const char *name)
create symlink between two objects.

**Parameters**

**struct kobject *kobj**
object whose directory we're creating the link in.

**struct kobject *target**
object we're pointing to.

**const char *name**
name of the symlink.

This function does the same as *sysfs_create_link()*, but it doesn't warn if the link already exists.

void **sysfs_remove_link**(struct kobject *kobj, const char *name)
remove symlink in object's directory.

**Parameters**

**struct kobject *kobj**
object we're acting for.

**const char *name**
name of the symlink to remove.

int **sysfs_rename_link_ns**(struct kobject *kobj, struct kobject *targ, const char *old, const char *new, const void *new_ns)
rename symlink in object's directory.

**Parameters**

**struct kobject *kobj**
object we're acting for.

**struct kobject *targ**
object we're pointing to.

**const char *old**
previous name of the symlink.

**const char *new**
new name of the symlink.

**const void *new_ns**
new namespace of the symlink.

A helper function for the common rename symlink idiom.

### 1.3.6 The debugfs filesystem

#### debugfs interface

struct dentry *__debugfs_lookup__(const char *name, struct dentry *parent)
look up an existing debugfs file

**Parameters**

**const char *name**
a pointer to a string containing the name of the file to look up.

**struct dentry *parent**
a pointer to the parent dentry of the file.

**Description**

This function will return a pointer to a dentry if it succeeds. If the file doesn't exist or an error occurs, `NULL` will be returned. The returned dentry must be passed to dput() when it is no longer needed.

If debugfs is not enabled in the kernel, the value `-ENODEV` will be returned.

struct dentry *__debugfs_create_file__(const char *name, umode_t mode, struct dentry *parent, void *data, const struct file_operations *fops)
create a file in the debugfs filesystem

**Parameters**

**const char *name**
a pointer to a string containing the name of the file to create.

**umode_t mode**
the permission that the file should have.

**struct dentry *parent**
a pointer to the parent dentry for this file. This should be a directory dentry if set. If this parameter is NULL, then the file will be created in the root of the debugfs filesystem.

**void *data**
a pointer to something that the caller will want to get to later on. The inode.i_private pointer will point to this value on the open() call.

**const struct file_operations *fops**
a pointer to a struct file_operations that should be used for this file.

**Description**

This is the basic "create a file" function for debugfs. It allows for a wide range of flexibility in creating a file, or a directory (if you want to create a directory, the *debugfs_create_dir()* function is recommended to be used instead.)

This function will return a pointer to a dentry if it succeeds. This pointer must be passed to the *debugfs_remove()* function when the file is to be removed (no automatic cleanup happens if your module is unloaded, you are responsible here.) If an error occurs, ERR_PTR(-ERROR) will be returned.

If debugfs is not enabled in the kernel, the value -ENODEV will be returned.

**NOTE**

it's expected that most callers should _ignore_ the errors returned by this function. Other debugfs functions handle the fact that the "dentry" passed to them could be an error and they don't crash in that case. Drivers should generally work fine even if debugfs fails to init anyway.

struct dentry ***debugfs_create_file_unsafe**(const char *name, umode_t mode, struct dentry *parent, void *data, const struct file_operations *fops)

create a file in the debugfs filesystem

**Parameters**

**const char *name**
a pointer to a string containing the name of the file to create.

**umode_t mode**
the permission that the file should have.

**struct dentry *parent**
a pointer to the parent dentry for this file. This should be a directory dentry if set. If this parameter is NULL, then the file will be created in the root of the debugfs filesystem.

**void *data**
a pointer to something that the caller will want to get to later on. The inode.i_private pointer will point to this value on the open() call.

**const struct file_operations *fops**
a pointer to a struct file_operations that should be used for this file.

**Description**

*debugfs_create_file_unsafe()* is completely analogous to *debugfs_create_file()*, the only difference being that the fops handed it will not get protected against file removals by the debugfs core.

It is your responsibility to protect your struct file_operation methods against file removals by means of *debugfs_file_get()* and *debugfs_file_put()*. ->open() is still protected by debugfs though.

Any struct file_operations defined by means of DEFINE_DEBUGFS_ATTRIBUTE() is protected against file removals and thus, may be used here.

void **debugfs_create_file_size**(const char *name, umode_t mode, struct dentry *parent, void *data, const struct file_operations *fops, loff_t file_size)

create a file in the debugfs filesystem

**Parameters**

**const char *name**
a pointer to a string containing the name of the file to create.

**umode_t mode**
the permission that the file should have.

**struct dentry *parent**
a pointer to the parent dentry for this file. This should be a directory dentry if set. If this parameter is NULL, then the file will be created in the root of the debugfs filesystem.

**void *data**
a pointer to something that the caller will want to get to later on. The inode.i_private pointer will point to this value on the open() call.

**const struct file_operations *fops**
a pointer to a struct file_operations that should be used for this file.

**loff_t file_size**
initial file size

**Description**

This is the basic "create a file" function for debugfs. It allows for a wide range of flexibility in creating a file, or a directory (if you want to create a directory, the *debugfs_create_dir()* function is recommended to be used instead.)

struct dentry *__debugfs_create_dir__(const char *name, struct dentry *parent)

create a directory in the debugfs filesystem

**Parameters**

**const char *name**
a pointer to a string containing the name of the directory to create.

**struct dentry *parent**
a pointer to the parent dentry for this file. This should be a directory dentry if set. If this parameter is NULL, then the directory will be created in the root of the debugfs filesystem.

**Description**

This function creates a directory in debugfs with the given name.

This function will return a pointer to a dentry if it succeeds. This pointer must be passed to the *debugfs_remove()* function when the file is to be removed (no automatic cleanup happens if your module is unloaded, you are responsible here.) If an error occurs, ERR_PTR(-ERROR) will be returned.

If debugfs is not enabled in the kernel, the value -ENODEV will be returned.

**NOTE**

it's expected that most callers should _ignore_ the errors returned by this function. Other debugfs functions handle the fact that the "dentry" passed to them could be an error and they don't crash in that case. Drivers should generally work fine even if debugfs fails to init anyway.

struct dentry *__debugfs_create_automount__(const char *name, struct dentry *parent, debugfs_automount_t f, void *data)

create automount point in the debugfs filesystem

**Parameters**

**const char *name**
a pointer to a string containing the name of the file to create.

**struct dentry *parent**
a pointer to the parent dentry for this file. This should be a directory dentry if set. If this parameter is NULL, then the file will be created in the root of the debugfs filesystem.

**debugfs_automount_t f**
function to be called when pathname resolution steps on that one.

**void *data**
opaque argument to pass to f().

**Description**

**f** should return what ->d_automount() would.

struct dentry *__debugfs_create_symlink__(const char *name, struct dentry *parent, const char *target)

create a symbolic link in the debugfs filesystem

**Parameters**

**const char *name**
a pointer to a string containing the name of the symbolic link to create.

**struct dentry *parent**
a pointer to the parent dentry for this symbolic link. This should be a directory dentry if set. If this parameter is NULL, then the symbolic link will be created in the root of the debugfs filesystem.

**const char *target**
a pointer to a string containing the path to the target of the symbolic link.

**Description**

This function creates a symbolic link with the given name in debugfs that links to the given target path.

This function will return a pointer to a dentry if it succeeds. This pointer must be passed to the *debugfs_remove()* function when the symbolic link is to be removed (no automatic cleanup happens if your module is unloaded, you are responsible here.) If an error occurs, ERR_PTR(-ERROR) will be returned.

If debugfs is not enabled in the kernel, the value -ENODEV will be returned.

void __debugfs_remove__(struct *dentry* *dentry)

recursively removes a directory

**Parameters**

**struct dentry *dentry**
a pointer to a the dentry of the directory to be removed. If this parameter is NULL or an error value, nothing will be done.

### Description

This function recursively removes a directory tree in debugfs that was previously created with a call to another debugfs function (like *debugfs_create_file()* or variants thereof.)

This function is required to be called in order for the file to be removed, no automatic cleanup of files will happen when a module is removed, you are responsible here.

void **debugfs_lookup_and_remove**(const char *name, struct dentry *parent)

lookup a directory or file and recursively remove it

### Parameters

**const char *name**
a pointer to a string containing the name of the item to look up.

**struct dentry *parent**
a pointer to the parent dentry of the item.

### Description

This is the equlivant of doing something like debugfs_remove(debugfs_lookup(..)) but with the proper reference counting handled for the directory being looked up.

struct dentry ***debugfs_rename**(struct dentry *old_dir, struct dentry *old_dentry, struct dentry *new_dir, const char *new_name)

rename a file/directory in the debugfs filesystem

### Parameters

**struct dentry *old_dir**
a pointer to the parent dentry for the renamed object. This should be a directory dentry.

**struct dentry *old_dentry**
dentry of an object to be renamed.

**struct dentry *new_dir**
a pointer to the parent dentry where the object should be moved. This should be a directory dentry.

**const char *new_name**
a pointer to a string containing the target name.

### Description

This function renames a file/directory in debugfs. The target must not exist for rename to succeed.

This function will return a pointer to old_dentry (which is updated to reflect renaming) if it succeeds. If an error occurs, ERR_PTR(-ERROR) will be returned.

If debugfs is not enabled in the kernel, the value `-ENODEV` will be returned.

bool **debugfs_initialized**(void)

Tells whether debugfs has been registered

### Parameters

**void**
no arguments

int **debugfs_file_get**(struct *dentry* *dentry)

mark the beginning of file data access

**Parameters**

**struct dentry *dentry**

the dentry object whose data is being accessed.

**Description**

Up to a matching call to *debugfs_file_put()*, any successive call into the file removing functions *debugfs_remove()* and debugfs_remove_recursive() will block. Since associated private file data may only get freed after a successful return of any of the removal functions, you may safely access it after a successful call to *debugfs_file_get()* without worrying about lifetime issues.

If -EIO is returned, the file has already been removed and thus, it is not safe to access any of its data. If, on the other hand, it is allowed to access the file data, zero is returned.

void **debugfs_file_put**(struct *dentry* *dentry)

mark the end of file data access

**Parameters**

**struct dentry *dentry**

the dentry object formerly passed to *debugfs_file_get()*.

**Description**

Allow any ongoing concurrent call into *debugfs_remove()* or debugfs_remove_recursive() blocked by a former call to *debugfs_file_get()* to proceed and return to its caller.

void **debugfs_enter_cancellation**(struct *file* *file, struct debugfs_cancellation *cancellation)

enter a debugfs cancellation

**Parameters**

**struct file *file**

the file being accessed

**struct debugfs_cancellation *cancellation**

the cancellation object, the cancel callback inside of it must be initialized

**Description**

When a debugfs file is removed it needs to wait for all active operations to complete. However, the operation itself may need to wait for hardware or completion of some asynchronous process or similar. As such, it may need to be cancelled to avoid long waits or even deadlocks.

This function can be used inside a debugfs handler that may need to be cancelled. As soon as this function is called, the cancellation's 'cancel' callback may be called, at which point the caller should proceed to call *debugfs_leave_cancellation()* and leave the debugfs handler function as soon as possible. Note that the 'cancel' callback is only ever called in the context of some kind of *debugfs_remove()*.

This function must be paired with *debugfs_leave_cancellation()*.

void **debugfs_leave_cancellation**(struct *file* *file, struct debugfs_cancellation *cancellation)

leave cancellation section

**Parameters**

**struct file *file**
the file being accessed

**struct debugfs_cancellation *cancellation**
the cancellation previously registered with *debugfs_enter_cancellation()*

**Description**

See the documentation of *debugfs_enter_cancellation()*.

void **debugfs_create_u8**(const char *name, umode_t mode, struct dentry *parent, u8 *value)

create a debugfs file that is used to read and write an unsigned 8-bit value

**Parameters**

**const char *name**
a pointer to a string containing the name of the file to create.

**umode_t mode**
the permission that the file should have

**struct dentry *parent**
a pointer to the parent dentry for this file. This should be a directory dentry if set. If this parameter is NULL, then the file will be created in the root of the debugfs filesystem.

**u8 *value**
a pointer to the variable that the file should read to and write from.

**Description**

This function creates a file in debugfs with the given name that contains the value of the variable **value**. If the **mode** variable is so set, it can be read from, and written to.

void **debugfs_create_u16**(const char *name, umode_t mode, struct dentry *parent, u16 *value)

create a debugfs file that is used to read and write an unsigned 16-bit value

**Parameters**

**const char *name**
a pointer to a string containing the name of the file to create.

**umode_t mode**
the permission that the file should have

**struct dentry *parent**
a pointer to the parent dentry for this file. This should be a directory dentry if set. If this parameter is NULL, then the file will be created in the root of the debugfs filesystem.

**u16 *value**
a pointer to the variable that the file should read to and write from.

**Description**

This function creates a file in debugfs with the given name that contains the value of the variable **value**. If the **mode** variable is so set, it can be read from, and written to.

void **debugfs_create_u32**(const char *name, umode_t mode, struct dentry *parent, u32 *value)

create a debugfs file that is used to read and write an unsigned 32-bit value

**Parameters**

**const char *name**
a pointer to a string containing the name of the file to create.

**umode_t mode**
the permission that the file should have

**struct dentry *parent**
a pointer to the parent dentry for this file. This should be a directory dentry if set. If this parameter is NULL, then the file will be created in the root of the debugfs filesystem.

**u32 *value**
a pointer to the variable that the file should read to and write from.

**Description**

This function creates a file in debugfs with the given name that contains the value of the variable **value**. If the **mode** variable is so set, it can be read from, and written to.

void **debugfs_create_u64**(const char *name, umode_t mode, struct dentry *parent, u64 *value)

create a debugfs file that is used to read and write an unsigned 64-bit value

**Parameters**

**const char *name**
a pointer to a string containing the name of the file to create.

**umode_t mode**
the permission that the file should have

**struct dentry *parent**
a pointer to the parent dentry for this file. This should be a directory dentry if set. If this parameter is NULL, then the file will be created in the root of the debugfs filesystem.

**u64 *value**
a pointer to the variable that the file should read to and write from.

**Description**

This function creates a file in debugfs with the given name that contains the value of the variable **value**. If the **mode** variable is so set, it can be read from, and written to.

void **debugfs_create_ulong**(const char *name, umode_t mode, struct dentry *parent, unsigned long *value)

create a debugfs file that is used to read and write an unsigned long value.

**Parameters**

**const char *name**
a pointer to a string containing the name of the file to create.

**umode_t mode**
the permission that the file should have

**struct dentry *parent**
a pointer to the parent dentry for this file. This should be a directory dentry if set. If this parameter is NULL, then the file will be created in the root of the debugfs filesystem.

**unsigned long *value**
a pointer to the variable that the file should read to and write from.

### Description

This function creates a file in debugfs with the given name that contains the value of the variable **value**. If the **mode** variable is so set, it can be read from, and written to.

void **debugfs_create_x8**(const char *name, umode_t mode, struct dentry *parent, u8 *value)
create a debugfs file that is used to read and write an unsigned 8-bit value

### Parameters

**const char *name**
a pointer to a string containing the name of the file to create.

**umode_t mode**
the permission that the file should have

**struct dentry *parent**
a pointer to the parent dentry for this file. This should be a directory dentry if set. If this parameter is NULL, then the file will be created in the root of the debugfs filesystem.

**u8 *value**
a pointer to the variable that the file should read to and write from.

void **debugfs_create_x16**(const char *name, umode_t mode, struct dentry *parent, u16 *value)
create a debugfs file that is used to read and write an unsigned 16-bit value

### Parameters

**const char *name**
a pointer to a string containing the name of the file to create.

**umode_t mode**
the permission that the file should have

**struct dentry *parent**
a pointer to the parent dentry for this file. This should be a directory dentry if set. If this parameter is NULL, then the file will be created in the root of the debugfs filesystem.

**u16 *value**
a pointer to the variable that the file should read to and write from.

void **debugfs_create_x32**(const char *name, umode_t mode, struct dentry *parent, u32 *value)
create a debugfs file that is used to read and write an unsigned 32-bit value

### Parameters

**const char *name**
a pointer to a string containing the name of the file to create.

**umode_t mode**
the permission that the file should have

**struct dentry *parent**
a pointer to the parent dentry for this file. This should be a directory dentry if set. If this parameter is NULL, then the file will be created in the root of the debugfs filesystem.

**u32 *value**
a pointer to the variable that the file should read to and write from.

void **debugfs_create_x64**(const char *name, umode_t mode, struct dentry *parent, u64 *value)

create a debugfs file that is used to read and write an unsigned 64-bit value

**Parameters**

**const char *name**
a pointer to a string containing the name of the file to create.

**umode_t mode**
the permission that the file should have

**struct dentry *parent**
a pointer to the parent dentry for this file. This should be a directory dentry if set. If this parameter is NULL, then the file will be created in the root of the debugfs filesystem.

**u64 *value**
a pointer to the variable that the file should read to and write from.

void **debugfs_create_size_t**(const char *name, umode_t mode, struct dentry *parent, size_t *value)

create a debugfs file that is used to read and write an size_t value

**Parameters**

**const char *name**
a pointer to a string containing the name of the file to create.

**umode_t mode**
the permission that the file should have

**struct dentry *parent**
a pointer to the parent dentry for this file. This should be a directory dentry if set. If this parameter is NULL, then the file will be created in the root of the debugfs filesystem.

**size_t *value**
a pointer to the variable that the file should read to and write from.

void **debugfs_create_atomic_t**(const char *name, umode_t mode, struct dentry *parent, atomic_t *value)

create a debugfs file that is used to read and write an atomic_t value

**Parameters**

**const char *name**
a pointer to a string containing the name of the file to create.

**umode_t mode**
the permission that the file should have

**struct dentry *parent**
: a pointer to the parent dentry for this file. This should be a directory dentry if set. If this parameter is NULL, then the file will be created in the root of the debugfs filesystem.

**atomic_t *value**
: a pointer to the variable that the file should read to and write from.

void **debugfs_create_bool**(const char *name, umode_t mode, struct dentry *parent, bool *value)

: create a debugfs file that is used to read and write a boolean value

**Parameters**

**const char *name**
: a pointer to a string containing the name of the file to create.

**umode_t mode**
: the permission that the file should have

**struct dentry *parent**
: a pointer to the parent dentry for this file. This should be a directory dentry if set. If this parameter is NULL, then the file will be created in the root of the debugfs filesystem.

**bool *value**
: a pointer to the variable that the file should read to and write from.

**Description**

This function creates a file in debugfs with the given name that contains the value of the variable **value**. If the **mode** variable is so set, it can be read from, and written to.

void **debugfs_create_str**(const char *name, umode_t mode, struct dentry *parent, char **value)

: create a debugfs file that is used to read and write a string value

**Parameters**

**const char *name**
: a pointer to a string containing the name of the file to create.

**umode_t mode**
: the permission that the file should have

**struct dentry *parent**
: a pointer to the parent dentry for this file. This should be a directory dentry if set. If this parameter is NULL, then the file will be created in the root of the debugfs filesystem.

**char **value**
: a pointer to the variable that the file should read to and write from.

**Description**

This function creates a file in debugfs with the given name that contains the value of the variable **value**. If the **mode** variable is so set, it can be read from, and written to.

struct dentry ***debugfs_create_blob**(const char *name, umode_t mode, struct dentry *parent, struct debugfs_blob_wrapper *blob)

: create a debugfs file that is used to read and write a binary blob

**Parameters**

**const char *name**
a pointer to a string containing the name of the file to create.

**umode_t mode**
the permission that the file should have

**struct dentry *parent**
a pointer to the parent dentry for this file. This should be a directory dentry if set. If this parameter is NULL, then the file will be created in the root of the debugfs filesystem.

**struct debugfs_blob_wrapper *blob**
a pointer to a struct debugfs_blob_wrapper which contains a pointer to the blob data and the size of the data.

**Description**

This function creates a file in debugfs with the given name that exports **blob->data** as a binary blob. If the **mode** variable is so set it can be read from and written to.

This function will return a pointer to a dentry if it succeeds. This pointer must be passed to the *debugfs_remove()* function when the file is to be removed (no automatic cleanup happens if your module is unloaded, you are responsible here.) If an error occurs, ERR_PTR(-ERROR) will be returned.

If debugfs is not enabled in the kernel, the value ERR_PTR(-ENODEV) will be returned.

void **debugfs_create_u32_array**(const char *name, umode_t mode, struct dentry *parent, struct debugfs_u32_array *array)

create a debugfs file that is used to read u32 array.

**Parameters**

**const char *name**
a pointer to a string containing the name of the file to create.

**umode_t mode**
the permission that the file should have.

**struct dentry *parent**
a pointer to the parent dentry for this file. This should be a directory dentry if set. If this parameter is NULL, then the file will be created in the root of the debugfs filesystem.

**struct debugfs_u32_array *array**
wrapper struct containing data pointer and size of the array.

**Description**

This function creates a file in debugfs with the given name that exports **array** as data. If the **mode** variable is so set it can be read from. Writing is not supported. Seek within the file is also not supported. Once array is created its size can not be changed.

void **debugfs_print_regs32**(struct seq_file *s, const struct debugfs_reg32 *regs, int nregs, void __iomem *base, char *prefix)

use seq_print to describe a set of registers

**Parameters**

**struct seq_file *s**
the seq_file structure being used to generate output

**const struct debugfs_reg32 *regs**
an array if struct debugfs_reg32 structures

**int nregs**
the length of the above array

**void __iomem *base**
the base address to be used in reading the registers

**char *prefix**
a string to be prefixed to every output line

**Description**

This function outputs a text block describing the current values of some 32-bit hardware registers. It is meant to be used within debugfs files based on seq_file that need to show registers, intermixed with other information. The prefix argument may be used to specify a leading string, because some peripherals have several blocks of identical registers, for example configuration of dma channels

void **debugfs_create_regset32**(const char *name, umode_t mode, struct dentry *parent, struct debugfs_regset32 *regset)

create a debugfs file that returns register values

**Parameters**

**const char *name**
a pointer to a string containing the name of the file to create.

**umode_t mode**
the permission that the file should have

**struct dentry *parent**
a pointer to the parent dentry for this file. This should be a directory dentry if set. If this parameter is `NULL`, then the file will be created in the root of the debugfs filesystem.

**struct debugfs_regset32 *regset**
a pointer to a struct debugfs_regset32, which contains a pointer to an array of register definitions, the array size and the base address where the register bank is to be found.

**Description**

This function creates a file in debugfs with the given name that reports the names and values of a set of 32-bit registers. If the **mode** variable is so set it can be read from. Writing is not supported.

void **debugfs_create_devm_seqfile**(struct device *dev, const char *name, struct dentry *parent, int (*read_fn)(struct seq_file *s, void *data))

create a debugfs file that is bound to device.

**Parameters**

**struct device *dev**
device related to this debugfs file.

**const char *name**
name of the debugfs file.

**struct dentry *parent**
: a pointer to the parent dentry for this file. This should be a directory dentry if set. If this parameter is NULL, then the file will be created in the root of the debugfs filesystem.

**int (*read_fn)(struct seq_file *s, void *data)**
: function pointer called to print the seq_file content.

## 1.4 splice and pipes

### 1.4.1 splice API

splice is a method for moving blocks of data around inside the kernel, without continually transferring them between the kernel and user space.

ssize_t **splice_to_pipe**(struct *pipe_inode_info* *pipe, struct splice_pipe_desc *spd)
: fill passed data into a pipe

**Parameters**

**struct pipe_inode_info *pipe**
: pipe to fill

**struct splice_pipe_desc *spd**
: data to fill

**Description**

> **spd** contains a map of pages and len/offset tuples, along with the struct pipe_buf_operations associated with these pages. This function will link that data to the pipe.

ssize_t **copy_splice_read**(struct file *in, loff_t *ppos, struct *pipe_inode_info* *pipe, size_t len, unsigned int flags)
: Copy data from a file and splice the copy into a pipe

**Parameters**

**struct file *in**
: The file to read from

**loff_t *ppos**
: Pointer to the file position to read from

**struct pipe_inode_info *pipe**
: The pipe to splice into

**size_t len**
: The amount to splice

**unsigned int flags**
: The SPLICE_F_* flags

**Description**

This function allocates a bunch of pages sufficient to hold the requested amount of data (but limited by the remaining pipe capacity), passes it to the file's ->read_iter() to read into and then splices the used pages into the pipe.

### Return

On success, the number of bytes read will be returned and ***ppos** will be updated if appropriate; 0 will be returned if there is no more data to be read; -EAGAIN will be returned if the pipe had no space, and some other negative error code will be returned on error. A short read may occur if the pipe has insufficient space, we reach the end of the data or we hit a hole.

int **splice_from_pipe_feed**(struct *pipe_inode_info* *pipe, struct splice_desc *sd, splice_actor *actor)

feed available data from a pipe to a file

### Parameters

**struct pipe_inode_info *pipe**
pipe to splice from

**struct splice_desc *sd**
information to **actor**

**splice_actor *actor**
handler that splices the data

### Description

This function loops over the pipe and calls **actor** to do the actual moving of a single *struct pipe_buffer* to the desired destination. It returns when there's no more buffers left in the pipe or if the requested number of bytes (**sd->total_len**) have been copied. It returns a positive number (one) if the pipe needs to be filled with more data, zero if the required number of bytes have been copied and -errno on error.

This, together with splice_from_pipe_{begin,end,next}, may be used to implement the functionality of *__splice_from_pipe()* when locking is required around copying the pipe buffers to the destination.

int **splice_from_pipe_next**(struct *pipe_inode_info* *pipe, struct splice_desc *sd)

wait for some data to splice from

### Parameters

**struct pipe_inode_info *pipe**
pipe to splice from

**struct splice_desc *sd**
information about the splice operation

### Description

This function will wait for some data and return a positive value (one) if pipe buffers are available. It will return zero or -errno if no more data needs to be spliced.

void **splice_from_pipe_begin**(struct splice_desc *sd)

start splicing from pipe

### Parameters

**struct splice_desc *sd**
information about the splice operation

### Description

This function should be called before a loop containing *splice_from_pipe_next()* and *splice_from_pipe_feed()* to initialize the necessary fields of **sd**.

void **splice_from_pipe_end**(struct *pipe_inode_info* *pipe, struct splice_desc *sd)

finish splicing from pipe

**Parameters**

**struct pipe_inode_info *pipe**
pipe to splice from

**struct splice_desc *sd**
information about the splice operation

**Description**

This function will wake up pipe writers if necessary. It should be called after a loop containing *splice_from_pipe_next()* and *splice_from_pipe_feed()*.

ssize_t **__splice_from_pipe**(struct *pipe_inode_info* *pipe, struct splice_desc *sd, splice_actor *actor)

splice data from a pipe to given actor

**Parameters**

**struct pipe_inode_info *pipe**
pipe to splice from

**struct splice_desc *sd**
information to **actor**

**splice_actor *actor**
handler that splices the data

**Description**

This function does little more than loop over the pipe and call **actor** to do the actual moving of a single *struct pipe_buffer* to the desired destination. See pipe_to_file, pipe_to_sendmsg, or pipe_to_user.

ssize_t **splice_from_pipe**(struct *pipe_inode_info* *pipe, struct file *out, loff_t *ppos, size_t len, unsigned int flags, splice_actor *actor)

splice data from a pipe to a file

**Parameters**

**struct pipe_inode_info *pipe**
pipe to splice from

**struct file *out**
file to splice to

**loff_t *ppos**
position in **out**

**size_t len**
how many bytes to splice

**unsigned int flags**
splice modifier flags

**splice_actor *actor**
handler that splices the data

**Description**

See __splice_from_pipe. This function locks the pipe inode, otherwise it's identical to *__splice_from_pipe()*.

ssize_t **iter_file_splice_write**(struct *pipe_inode_info* *pipe, struct file *out, loff_t *ppos, size_t len, unsigned int flags)

splice data from a pipe to a file

**Parameters**

**struct pipe_inode_info *pipe**
pipe info

**struct file *out**
file to write to

**loff_t *ppos**
position in **out**

**size_t len**
number of bytes to splice

**unsigned int flags**
splice modifier flags

**Description**

Will either move or copy pages (determined by **flags** options) from the given pipe inode to the given file. This one is ->write_iter-based.

ssize_t **splice_to_socket**(struct *pipe_inode_info* *pipe, struct file *out, loff_t *ppos, size_t len, unsigned int flags)

splice data from a pipe to a socket

**Parameters**

**struct pipe_inode_info *pipe**
pipe to splice from

**struct file *out**
socket to write to

**loff_t *ppos**
position in **out**

**size_t len**
number of bytes to splice

**unsigned int flags**
splice modifier flags

**Description**

Will send **len** bytes from the pipe to a network socket. No data copying is involved.

ssize_t **vfs_splice_read**(struct file *in, loff_t *ppos, struct *pipe_inode_info* *pipe, size_t len, unsigned int flags)

Read data from a file and splice it into a pipe

**Parameters**

**struct file *in**
File to splice from

**loff_t *ppos**
Input file offset

**struct pipe_inode_info *pipe**
Pipe to splice to

**size_t len**
Number of bytes to splice

**unsigned int flags**
Splice modifier flags (SPLICE_F_*)

**Description**

Splice the requested amount of data from the input file to the pipe. This is synchronous as the caller must hold the pipe lock across the entire operation.

If successful, it returns the amount of data spliced, 0 if it hit the EOF or a hole and a negative error code otherwise.

ssize_t **splice_direct_to_actor**(struct file *in, struct splice_desc *sd, splice_direct_actor *actor)

splices data directly between two non-pipes

**Parameters**

**struct file *in**
file to splice from

**struct splice_desc *sd**
actor information on where to splice to

**splice_direct_actor *actor**
handles the data splicing

**Description**

This is a special case helper to splice directly between two points, without requiring an explicit pipe. Internally an allocated pipe is cached in the process, and reused during the lifetime of that process.

ssize_t **do_splice_direct**(struct file *in, loff_t *ppos, struct file *out, loff_t *opos, size_t len, unsigned int flags)

splices data directly between two files

**Parameters**

**struct file *in**
file to splice from

**loff_t *ppos**
input file offset

**struct file *out**
file to splice to

**loff_t *opos**
output file offset

**size_t len**
number of bytes to splice

**unsigned int flags**
splice modifier flags

**Description**

For use by do_sendfile(). splice can easily emulate sendfile, but doing it in the application would incur an extra system call (splice in + splice out, as compared to just sendfile()). So this helper can splice directly through a process-private pipe.

Callers already called rw_verify_area() on the entire range.

ssize_t **splice_file_range**(struct file *in, loff_t *ppos, struct file *out, loff_t *opos, size_t len)
splices data between two files for copy_file_range()

**Parameters**

**struct file *in**
file to splice from

**loff_t *ppos**
input file offset

**struct file *out**
file to splice to

**loff_t *opos**
output file offset

**size_t len**
number of bytes to splice

**Description**

For use by ->copy_file_range() methods. Like *do_splice_direct()*, but vfs_copy_file_range() already holds start_file_write() on **out** file.

Callers already called rw_verify_area() on the entire range.

### 1.4.2 pipes API

Pipe interfaces are all for in-kernel (builtin image) use. They are not exported for use by modules.

struct **pipe_buffer**
: a linux kernel pipe buffer

**Definition**:

```
struct pipe_buffer {
    struct page *page;
    unsigned int offset, len;
    const struct pipe_buf_operations *ops;
    unsigned int flags;
    unsigned long private;
};
```

**Members**

**page**
: the page containing the data for the pipe buffer

**offset**
: offset of data inside the **page**

**len**
: length of data inside the **page**

**ops**
: operations associated with this buffer. See **pipe_buf_operations**.

**flags**
: pipe buffer flags. See above.

**private**
: private data owned by the ops.

struct **pipe_inode_info**
: a linux kernel pipe

**Definition**:

```
struct pipe_inode_info {
    struct mutex mutex;
    wait_queue_head_t rd_wait, wr_wait;
    unsigned int head;
    unsigned int tail;
    unsigned int max_usage;
    unsigned int ring_size;
    unsigned int nr_accounted;
    unsigned int readers;
    unsigned int writers;
    unsigned int files;
    unsigned int r_counter;
    unsigned int w_counter;
```

```
    bool poll_usage;
#ifdef CONFIG_WATCH_QUEUE;
    bool note_loss;
#endif;
    struct page *tmp_page;
    struct fasync_struct *fasync_readers;
    struct fasync_struct *fasync_writers;
    struct pipe_buffer *bufs;
    struct user_struct *user;
#ifdef CONFIG_WATCH_QUEUE;
    struct watch_queue *watch_queue;
#endif;
};
```

### Members

**mutex**
mutex protecting the whole thing

**rd_wait**
reader wait point in case of empty pipe

**wr_wait**
writer wait point in case of full pipe

**head**
The point of buffer production

**tail**
The point of buffer consumption

**max_usage**
The maximum number of slots that may be used in the ring

**ring_size**
total number of buffers (should be a power of 2)

**nr_accounted**
The amount this pipe accounts for in user->pipe_bufs

**readers**
number of current readers of this pipe

**writers**
number of current writers of this pipe

**files**
number of struct file referring this pipe (protected by ->i_lock)

**r_counter**
reader counter

**w_counter**
writer counter

**poll_usage**
is this pipe used for epoll, which has crazy wakeups?

**note_loss**
The next read() should insert a data-lost message

**tmp_page**
cached released page

**fasync_readers**
reader side fasync

**fasync_writers**
writer side fasync

**bufs**
the circular array of pipe buffers

**user**
the user who created this pipe

**watch_queue**
If this pipe is a watch_queue, this is the stuff for that

bool **pipe_has_watch_queue**(const struct *pipe_inode_info* *pipe)
Check whether the pipe is a watch_queue, i.e. it was created with O_NOTIFICATION_PIPE

**Parameters**

**const struct pipe_inode_info *pipe**
The pipe to check

**Return**

true if pipe is a watch queue, false otherwise.

bool **pipe_empty**(unsigned int head, unsigned int tail)
Return true if the pipe is empty

**Parameters**

**unsigned int head**
The pipe ring head pointer

**unsigned int tail**
The pipe ring tail pointer

unsigned int **pipe_occupancy**(unsigned int head, unsigned int tail)
Return number of slots used in the pipe

**Parameters**

**unsigned int head**
The pipe ring head pointer

**unsigned int tail**
The pipe ring tail pointer

bool **pipe_full**(unsigned int head, unsigned int tail, unsigned int limit)
Return true if the pipe is full

**Parameters**

**unsigned int head**
The pipe ring head pointer

**unsigned int tail**
The pipe ring tail pointer

**unsigned int limit**
The maximum amount of slots available.

struct *pipe_buffer* ***pipe_buf**(const struct *pipe_inode_info* *pipe, unsigned int slot)
Return the pipe buffer for the specified slot in the pipe ring

**Parameters**

**const struct pipe_inode_info *pipe**
The pipe to access

**unsigned int slot**
The slot of interest

struct *pipe_buffer* ***pipe_head_buf**(const struct *pipe_inode_info* *pipe)
Return the pipe buffer at the head of the pipe ring

**Parameters**

**const struct pipe_inode_info *pipe**
The pipe to access

bool **pipe_buf_get**(struct *pipe_inode_info* *pipe, struct *pipe_buffer* *buf)
get a reference to a pipe_buffer

**Parameters**

**struct pipe_inode_info *pipe**
the pipe that the buffer belongs to

**struct pipe_buffer *buf**
the buffer to get a reference to

**Return**

`true` if the reference was successfully obtained.

void **pipe_buf_release**(struct *pipe_inode_info* *pipe, struct *pipe_buffer* *buf)
put a reference to a pipe_buffer

**Parameters**

**struct pipe_inode_info *pipe**
the pipe that the buffer belongs to

**struct pipe_buffer *buf**
the buffer to put a reference to

int **pipe_buf_confirm**(struct *pipe_inode_info* *pipe, struct *pipe_buffer* *buf)
verify contents of the pipe buffer

**Parameters**

**struct pipe_inode_info *pipe**
the pipe that the buffer belongs to

**struct pipe_buffer *buf**
the buffer to confirm

bool **pipe_buf_try_steal**(struct *pipe_inode_info* *pipe, struct *pipe_buffer* *buf)
attempt to take ownership of a pipe_buffer

**Parameters**

**struct pipe_inode_info *pipe**
the pipe that the buffer belongs to

**struct pipe_buffer *buf**
the buffer to attempt to steal

bool **generic_pipe_buf_try_steal**(struct *pipe_inode_info* *pipe, struct *pipe_buffer* *buf)
attempt to take ownership of a *pipe_buffer*

**Parameters**

**struct pipe_inode_info *pipe**
the pipe that the buffer belongs to

**struct pipe_buffer *buf**
the buffer to attempt to steal

**Description**

This function attempts to steal the `struct page` attached to **buf**. If successful, this function returns 0 and returns with the page locked. The caller may then reuse the page for whatever he wishes; the typical use is insertion into a different file page cache.

bool **generic_pipe_buf_get**(struct *pipe_inode_info* *pipe, struct *pipe_buffer* *buf)
get a reference to a *struct pipe_buffer*

**Parameters**

**struct pipe_inode_info *pipe**
the pipe that the buffer belongs to

**struct pipe_buffer *buf**
the buffer to get a reference to

**Description**

This function grabs an extra reference to **buf**. It's used in the tee() system call, when we duplicate the buffers in one pipe into another.

void **generic_pipe_buf_release**(struct *pipe_inode_info* *pipe, struct *pipe_buffer* *buf)
put a reference to a *struct pipe_buffer*

**Parameters**

**struct pipe_inode_info *pipe**
the pipe that the buffer belongs to

**struct pipe_buffer *buf**
the buffer to put a reference to

**Description**

This function releases a reference to **buf**.

## 1.5 Locking

The text below describes the locking rules for VFS-related methods. It is (believed to be) up-to-date. *Please,* if you change anything in prototypes or locking protocols - update this file. And update the relevant instances in the tree, don't leave that to maintainers of filesystems/devices/ etc. At the very least, put the list of dubious cases in the end of this file. Don't turn it into log - maintainers of out-of-the-tree code are supposed to be able to use diff(1).

Thing currently missing here: socket operations. Alexey?

### 1.5.1 dentry_operations

prototypes:

```
int (*d_revalidate)(struct dentry *, unsigned int);
int (*d_weak_revalidate)(struct dentry *, unsigned int);
int (*d_hash)(const struct dentry *, struct qstr *);
int (*d_compare)(const struct dentry *,
                 unsigned int, const char *, const struct qstr *);
int (*d_delete)(struct dentry *);
int (*d_init)(struct dentry *);
void (*d_release)(struct dentry *);
void (*d_iput)(struct dentry *, struct inode *);
char *(*d_dname)((struct dentry *dentry, char *buffer, int buflen);
struct vfsmount *(*d_automount)(struct path *path);
int (*d_manage)(const struct path *, bool);
struct dentry *(*d_real)(struct dentry *, const struct inode *);
```

locking rules:

| ops | rename_lock | ->d_lock | may block | rcu-walk |
|---|---|---|---|---|
| d_revalidate: | no | no | yes (ref-walk) | maybe |
| d_weak_revalidate: | no | no | yes | no |
| d_hash | no | no | no | maybe |
| d_compare: | yes | no | no | maybe |
| d_delete: | no | yes | no | no |
| d_init: | no | no | yes | no |
| d_release: | no | no | yes | no |
| d_prune: | no | yes | no | no |
| d_iput: | no | no | yes | no |
| d_dname: | no | no | no | no |
| d_automount: | no | no | yes | no |
| d_manage: | no | no | yes (ref-walk) | maybe |
| d_real | no | no | yes | no |

### 1.5.2 inode_operations

prototypes:

```
int (*create) (struct mnt_idmap *, struct inode *,struct dentry *,umode_t,␣
↪bool);
struct dentry * (*lookup) (struct inode *,struct dentry *, unsigned int);
int (*link) (struct dentry *,struct inode *,struct dentry *);
int (*unlink) (struct inode *,struct dentry *);
int (*symlink) (struct mnt_idmap *, struct inode *,struct dentry *,const char␣
↪*);
int (*mkdir) (struct mnt_idmap *, struct inode *,struct dentry *,umode_t);
int (*rmdir) (struct inode *,struct dentry *);
int (*mknod) (struct mnt_idmap *, struct inode *,struct dentry *,umode_t,dev_
↪t);
int (*rename) (struct mnt_idmap *, struct inode *, struct dentry *,
                struct inode *, struct dentry *, unsigned int);
int (*readlink) (struct dentry *, char __user *,int);
const char *(*get_link) (struct dentry *, struct inode *, struct delayed_call␣
↪*);
void (*truncate) (struct inode *);
int (*permission) (struct mnt_idmap *, struct inode *, int, unsigned int);
struct posix_acl * (*get_inode_acl)(struct inode *, int, bool);
int (*setattr) (struct mnt_idmap *, struct dentry *, struct iattr *);
int (*getattr) (struct mnt_idmap *, const struct path *, struct kstat *, u32,␣
↪unsigned int);
ssize_t (*listxattr) (struct dentry *, char *, size_t);
int (*fiemap)(struct inode *, struct fiemap_extent_info *, u64 start, u64 len);
void (*update_time)(struct inode *, struct timespec *, int);
int (*atomic_open)(struct inode *, struct dentry *,
                       struct file *, unsigned open_flag,
                       umode_t create_mode);
int (*tmpfile) (struct mnt_idmap *, struct inode *,
                struct file *, umode_t);
int (*fileattr_set)(struct mnt_idmap *idmap,
                    struct dentry *dentry, struct fileattr *fa);
int (*fileattr_get)(struct dentry *dentry, struct fileattr *fa);
struct posix_acl * (*get_acl)(struct mnt_idmap *, struct dentry *, int);
struct offset_ctx *(*get_offset_ctx)(struct inode *inode);
```

**locking rules:**

all may block

| ops | i_rwsem(inode) |
|---|---|
| lookup: | shared |
| create: | exclusive |
| link: | exclusive (both) |
| mknod: | exclusive |
| symlink: | exclusive |
| mkdir: | exclusive |
| unlink: | exclusive (both) |
| rmdir: | exclusive (both)(see below) |
| rename: | exclusive (both parents, some children) (see below) |
| readlink: | no |
| get_link: | no |
| setattr: | exclusive |
| permission: | no (may not block if called in rcu-walk mode) |
| get_inode_acl: | no |
| get_acl: | no |
| getattr: | no |
| listxattr: | no |
| fiemap: | no |
| update_time: | no |
| atomic_open: | shared (exclusive if O_CREAT is set in open flags) |
| tmpfile: | no |
| fileattr_get: | no or exclusive |
| fileattr_set: | exclusive |
| get_offset_ctx | no |

Additionally, ->rmdir(), ->unlink() and ->rename() have ->i_rwsem exclusive on victim. cross-directory ->rename() has (per-superblock) ->s_vfs_rename_sem. ->unlink() and ->rename() have ->i_rwsem exclusive on all non-directories involved. ->rename() has ->i_rwsem exclusive on any subdirectory that changes parent.

See *Directory Locking* for more detailed discussion of the locking scheme for directory operations.

### 1.5.3 xattr_handler operations

prototypes:

```
bool (*list)(struct dentry *dentry);
int (*get)(const struct xattr_handler *handler, struct dentry *dentry,
           struct inode *inode, const char *name, void *buffer,
           size_t size);
int (*set)(const struct xattr_handler *handler,
           struct mnt_idmap *idmap,
           struct dentry *dentry, struct inode *inode, const char *name,
           const void *buffer, size_t size, int flags);
```

**locking rules:**
all may block

| ops | i_rwsem(inode) |
|---|---|
| list: | no |
| get: | no |
| set: | exclusive |

### 1.5.4 super_operations

prototypes:

```
struct inode *(*alloc_inode)(struct super_block *sb);
void (*free_inode)(struct inode *);
void (*destroy_inode)(struct inode *);
void (*dirty_inode) (struct inode *, int flags);
int (*write_inode) (struct inode *, struct writeback_control *wbc);
int (*drop_inode) (struct inode *);
void (*evict_inode) (struct inode *);
void (*put_super) (struct super_block *);
int (*sync_fs)(struct super_block *sb, int wait);
int (*freeze_fs) (struct super_block *);
int (*unfreeze_fs) (struct super_block *);
int (*statfs) (struct dentry *, struct kstatfs *);
int (*remount_fs) (struct super_block *, int *, char *);
void (*umount_begin) (struct super_block *);
int (*show_options)(struct seq_file *, struct dentry *);
ssize_t (*quota_read)(struct super_block *, int, char *, size_t, loff_t);
ssize_t (*quota_write)(struct super_block *, int, const char *, size_t, loff_
↪t);
```

**locking rules:**

All may block [not true, see below]

| ops | s_umount | note |
|---|---|---|
| alloc_inode: | | |
| free_inode: | | called from RCU callback |
| destroy_inode: | | |
| dirty_inode: | | |
| write_inode: | | |
| drop_inode: | | !!!inode->i_lock!!! |
| evict_inode: | | |
| put_super: | write | |
| sync_fs: | read | |
| freeze_fs: | write | |
| unfreeze_fs: | write | |
| statfs: | maybe(read) | (see below) |
| remount_fs: | write | |
| umount_begin: | no | |
| show_options: | no | (namespace_sem) |
| quota_read: | no | (see below) |
| quota_write: | no | (see below) |

->statfs() has s_umount (shared) when called by ustat(2) (native or compat), but that's an accident of bad API; s_umount is used to pin the superblock down when we only have dev_t given us by userland to identify the superblock. Everything else (statfs(), fstatfs(), etc.) doesn't hold it when calling ->statfs() - superblock is pinned down by resolving the pathname passed to syscall.

->quota_read() and ->quota_write() functions are both guaranteed to be the only ones operating on the quota file by the quota code (via dqio_sem) (unless an admin really wants to screw up something and writes to quota files with quotas on). For other details about locking see also dquot_operations section.

### 1.5.5 file_system_type

prototypes:

```
struct dentry *(*mount) (struct file_system_type *, int,
               const char *, void *);
void (*kill_sb) (struct super_block *);
```

locking rules:

| ops | may block |
|---|---|
| mount | yes |
| kill_sb | yes |

->mount() returns ERR_PTR or the root dentry; its superblock should be locked on return.

->kill_sb() takes a write-locked superblock, does all shutdown work on it, unlocks and drops the reference.

### 1.5.6 address_space_operations

prototypes:

```
int (*writepage)(struct page *page, struct writeback_control *wbc);
int (*read_folio)(struct file *, struct folio *);
int (*writepages)(struct address_space *, struct writeback_control *);
bool (*dirty_folio)(struct address_space *, struct folio *folio);
void (*readahead)(struct readahead_control *);
int (*write_begin)(struct file *, struct address_space *mapping,
                        loff_t pos, unsigned len,
                        struct page **pagep, void **fsdata);
int (*write_end)(struct file *, struct address_space *mapping,
                        loff_t pos, unsigned len, unsigned copied,
                        struct page *page, void *fsdata);
sector_t (*bmap)(struct address_space *, sector_t);
void (*invalidate_folio) (struct folio *, size_t start, size_t len);
bool (*release_folio)(struct folio *, gfp_t);
void (*free_folio)(struct folio *);
int (*direct_IO)(struct kiocb *, struct iov_iter *iter);
int (*migrate_folio)(struct address_space *, struct folio *dst,
                struct folio *src, enum migrate_mode);
```

```
int (*launder_folio)(struct folio *);
bool (*is_partially_uptodate)(struct folio *, size_t from, size_t count);
int (*error_remove_folio)(struct address_space *, struct folio *);
int (*swap_activate)(struct swap_info_struct *sis, struct file *f, sector_t␣
↪*span)
int (*swap_deactivate)(struct file *);
int (*swap_rw)(struct kiocb *iocb, struct iov_iter *iter);
```

**locking rules:**

All except dirty_folio and free_folio may block

| ops | folio locked | i_rwsem | invalidate_lock |
|---|---|---|---|
| writepage: | yes, unlocks (see below) | | |
| read_folio: | yes, unlocks | | shared |
| writepages: | | | |
| dirty_folio: | maybe | | |
| readahead: | yes, unlocks | | shared |
| write_begin: | locks the page | exclusive | |
| write_end: | yes, unlocks | exclusive | |
| bmap: | | | |
| invalidate_folio: | yes | | exclusive |
| release_folio: | yes | | |
| free_folio: | yes | | |
| direct_IO: | | | |
| migrate_folio: | yes (both) | | |
| launder_folio: | yes | | |
| is_partially_uptodate: | yes | | |
| error_remove_folio: | yes | | |
| swap_activate: | no | | |
| swap_deactivate: | no | | |
| swap_rw: | yes, unlocks | | |

->write_begin(), ->write_end() and ->read_folio() may be called from the request handler (/dev/loop).

->read_folio() unlocks the folio, either synchronously or via I/O completion.

->readahead() unlocks the folios that I/O is attempted on like ->read_folio().

->writepage() is used for two purposes: for ”memory cleansing” and for ”sync”. These are quite different operations and the behaviour may differ depending upon the mode.

If writepage is called for sync (wbc->sync_mode != WBC_SYNC_NONE) then it *must* start I/O against the page, even if that would involve blocking on in-progress I/O.

If writepage is called for memory cleansing (sync_mode == WBC_SYNC_NONE) then its role is to get as much writeout underway as possible. So writepage should try to avoid blocking against currently-in-progress I/O.

If the filesystem is not called for ”sync” and it determines that it would need to block against in-progress I/O to be able to start new I/O against the page the filesystem should redirty the page with redirty_page_for_writepage(), then unlock the page and return zero. This may also be done to avoid internal deadlocks, but rarely.

If the filesystem is called for sync then it must wait on any in-progress I/O and then start new I/O.

The filesystem should unlock the page synchronously, before returning to the caller, unless ->writepage() returns special WRITEPAGE_ACTIVATE value. WRITEPAGE_ACTIVATE means that page cannot really be written out currently, and VM should stop calling ->writepage() on this page for some time. VM does this by moving page to the head of the active list, hence the name.

Unless the filesystem is going to redirty_page_for_writepage(), unlock the page and return zero, writepage *must* run set_page_writeback() against the page, followed by unlocking it. Once set_page_writeback() has been run against the page, write I/O can be submitted and the write I/O completion handler must run end_page_writeback() once the I/O is complete. If no I/O is submitted, the filesystem must run end_page_writeback() against the page before returning from writepage.

That is: after 2.5.12, pages which are under writeout are *not* locked. Note, if the filesystem needs the page to be locked during writeout, that is ok, too, the page is allowed to be unlocked at any point in time between the calls to set_page_writeback() and end_page_writeback().

Note, failure to run either redirty_page_for_writepage() or the combination of set_page_writeback()/end_page_writeback() on a page submitted to writepage will leave the page itself marked clean but it will be tagged as dirty in the radix tree. This incoherency can lead to all sorts of hard-to-debug problems in the filesystem like having dirty inodes at umount and losing written data.

->writepages() is used for periodic writeback and for syscall-initiated sync operations. The address_space should start I/O against at least `*nr_to_write` pages. `*nr_to_write` must be decremented for each page which is written. The address_space implementation may write more (or less) pages than `*nr_to_write` asks for, but it should try to be reasonably close. If nr_to_write is NULL, all dirty pages must be written.

writepages should _only_ write pages which are present on mapping->io_pages.

->dirty_folio() is called from various places in the kernel when the target folio is marked as needing writeback. The folio cannot be truncated because either the caller holds the folio lock, or the caller has found the folio while holding the page table lock which will block truncation.

->*bmap()* is currently used by legacy ioctl() (FIBMAP) provided by some filesystems and by the swapper. The latter will eventually go away. Please, keep it that way and don't breed new callers.

->invalidate_folio() is called when the filesystem must attempt to drop some or all of the buffers from the page when it is being truncated. It returns zero on success. The filesystem must exclusively acquire invalidate_lock before invalidating page cache in truncate / hole punch path (and thus calling into ->invalidate_folio) to block races between page cache invalidation and page cache filling functions (fault, read, ...).

->release_folio() is called when the MM wants to make a change to the folio that would invalidate the filesystem's private data. For example, it may be about to be removed from the address_space or split. The folio is locked and not under writeback. It may be dirty. The gfp parameter is not usually used for allocation, but rather to indicate what the filesystem may do to attempt to free the private data. The filesystem may return false to indicate that the folio's private data cannot be freed. If it returns true, it should have already removed the private data from the folio. If a filesystem does not provide a ->release_folio method, the pagecache will assume that private data is buffer_heads and call try_to_free_buffers().

->free_folio() is called when the kernel has dropped the folio from the page cache.

->launder_folio() may be called prior to releasing a folio if it is still found to be dirty. It returns zero if the folio was successfully cleaned, or an error value if not. Note that in order to prevent the folio getting mapped back in and redirtied, it needs to be kept locked across the entire operation.

->swap_activate() will be called to prepare the given file for swap. It should perform any validation and preparation necessary to ensure that writes can be performed with minimal memory allocation. It should call add_swap_extent(), or the helper iomap_swapfile_activate(), and return the number of extents added. If IO should be submitted through ->swap_rw(), it should set SWP_FS_OPS, otherwise IO will be submitted directly to the block device `sis->bdev`.

->swap_deactivate() will be called in the sys_swapoff() path after ->swap_activate() returned success.

->swap_rw will be called for swap IO if SWP_FS_OPS was set by ->swap_activate().

### 1.5.7 file_lock_operations

prototypes:

```
void (*fl_copy_lock)(struct file_lock *, struct file_lock *);
void (*fl_release_private)(struct file_lock *);
```

locking rules:

| ops | inode->i_lock | may block |
|---|---|---|
| fl_copy_lock: | yes | no |
| fl_release_private: | maybe | maybe[1]_ |

### 1.5.8 lock_manager_operations

prototypes:

```
void (*lm_notify)(struct file_lock *);  /* unblock callback */
int (*lm_grant)(struct file_lock *, struct file_lock *, int);
void (*lm_break)(struct file_lock *); /* break_lease callback */
int (*lm_change)(struct file_lock **, int);
bool (*lm_breaker_owns_lease)(struct file_lock *);
bool (*lm_lock_expirable)(struct file_lock *);
void (*lm_expire_lock)(void);
```

locking rules:

| ops | flc_lock | blocked_lock_lock | may block |
|---|---|---|---|
| lm_notify: | no | yes | no |
| lm_grant: | no | no | no |
| lm_break: | yes | no | no |
| lm_change | yes | no | no |
| lm_breaker_owns_lease: | yes | no | no |
| lm_lock_expirable | yes | no | no |
| lm_expire_lock | no | no | yes |

### 1.5.9 buffer_head

prototypes:

```
void (*b_end_io)(struct buffer_head *bh, int uptodate);
```

locking rules:

called from interrupts. In other words, extreme care is needed here. bh is locked, but that's all warranties we have here. Currently only RAID1, highmem, fs/buffer.c, and fs/ntfs/aops.c are providing these. Block devices call this method upon the IO completion.

### 1.5.10 block_device_operations

prototypes:

```
int (*open) (struct block_device *, fmode_t);
int (*release) (struct gendisk *, fmode_t);
int (*ioctl) (struct block_device *, fmode_t, unsigned, unsigned long);
int (*compat_ioctl) (struct block_device *, fmode_t, unsigned, unsigned long);
int (*direct_access) (struct block_device *, sector_t, void **,
                      unsigned long *);
void (*unlock_native_capacity) (struct gendisk *);
int (*getgeo)(struct block_device *, struct hd_geometry *);
void (*swap_slot_free_notify) (struct block_device *, unsigned long);
```

locking rules:

| ops | open_mutex |
|---|---|
| open: | yes |
| release: | yes |
| ioctl: | no |
| compat_ioctl: | no |
| direct_access: | no |
| unlock_native_capacity: | no |
| getgeo: | no |
| swap_slot_free_notify: | no (see below) |

swap_slot_free_notify is called with swap_lock and sometimes the page lock held.

### 1.5.11 file_operations

prototypes:

```
loff_t (*llseek) (struct file *, loff_t, int);
ssize_t (*read) (struct file *, char __user *, size_t, loff_t *);
ssize_t (*write) (struct file *, const char __user *, size_t, loff_t *);
ssize_t (*read_iter) (struct kiocb *, struct iov_iter *);
ssize_t (*write_iter) (struct kiocb *, struct iov_iter *);
int (*iopoll) (struct kiocb *kiocb, bool spin);
int (*iterate_shared) (struct file *, struct dir_context *);
__poll_t (*poll) (struct file *, struct poll_table_struct *);
long (*unlocked_ioctl) (struct file *, unsigned int, unsigned long);
long (*compat_ioctl) (struct file *, unsigned int, unsigned long);
int (*mmap) (struct file *, struct vm_area_struct *);
int (*open) (struct inode *, struct file *);
int (*flush) (struct file *);
int (*release) (struct inode *, struct file *);
int (*fsync) (struct file *, loff_t start, loff_t end, int datasync);
int (*fasync) (int, struct file *, int);
int (*lock) (struct file *, int, struct file_lock *);
unsigned long (*get_unmapped_area)(struct file *, unsigned long,
                unsigned long, unsigned long, unsigned long);
int (*check_flags)(int);
int (*flock) (struct file *, int, struct file_lock *);
ssize_t (*splice_write)(struct pipe_inode_info *, struct file *, loff_t *,
                size_t, unsigned int);
ssize_t (*splice_read)(struct file *, loff_t *, struct pipe_inode_info *,
                size_t, unsigned int);
int (*setlease)(struct file *, long, struct file_lock **, void **);
long (*fallocate)(struct file *, int, loff_t, loff_t);
void (*show_fdinfo)(struct seq_file *m, struct file *f);
unsigned (*mmap_capabilities)(struct file *);
ssize_t (*copy_file_range)(struct file *, loff_t, struct file *,
                loff_t, size_t, unsigned int);
loff_t (*remap_file_range)(struct file *file_in, loff_t pos_in,
                struct file *file_out, loff_t pos_out,
                loff_t len, unsigned int remap_flags);
int (*fadvise)(struct file *, loff_t, loff_t, int);
```

**locking rules:**

All may block.

->llseek() locking has moved from llseek to the individual llseek implementations. If your fs is not using generic_file_llseek, you need to acquire and release the appropriate locks in your ->llseek(). For many filesystems, it is probably safe to acquire the inode mutex or just to use i_size_read() instead. Note: this does not protect the file->f_pos against concurrent modifications since this is something the userspace has to take care about.

->iterate_shared() is called with i_rwsem held for reading, and with the file f_pos_lock held exclusively

->fasync() is responsible for maintaining the FASYNC bit in filp->f_flags. Most instances call

fasync_helper(), which does that maintenance, so it's not normally something one needs to worry about. Return values > 0 will be mapped to zero in the VFS layer.

->readdir() and ->ioctl() on directories must be changed. Ideally we would move ->readdir() to inode_operations and use a separate method for directory ->ioctl() or kill the latter completely. One of the problems is that for anything that resembles union-mount we won't have a struct file for all components. And there are other reasons why the current interface is a mess...

->read on directories probably must go away - we should just enforce -EISDIR in sys_read() and friends.

->setlease operations should call *generic_setlease()* before or after setting the lease within the individual filesystem to record the result of the operation

->fallocate implementation must be really careful to maintain page cache consistency when punching holes or performing other operations that invalidate page cache contents. Usually the filesystem needs to call truncate_inode_pages_range() to invalidate relevant range of the page cache. However the filesystem usually also needs to update its internal (and on disk) view of file offset -> disk block mapping. Until this update is finished, the filesystem needs to block page faults and reads from reloading now-stale page cache contents from the disk. Since VFS acquires mapping->invalidate_lock in shared mode when loading pages from disk (filemap_fault(), filemap_read(), readahead paths), the fallocate implementation must take the invalidate_lock to prevent reloading.

->copy_file_range and ->remap_file_range implementations need to serialize against modifications of file data while the operation is running. For blocking changes through write(2) and similar operations inode->i_rwsem can be used. To block changes to file contents via a memory mapping during the operation, the filesystem must take mapping->invalidate_lock to coordinate with ->page_mkwrite.

### 1.5.12 dquot_operations

prototypes:

```
int (*write_dquot) (struct dquot *);
int (*acquire_dquot) (struct dquot *);
int (*release_dquot) (struct dquot *);
int (*mark_dirty) (struct dquot *);
int (*write_info) (struct super_block *, int);
```

These operations are intended to be more or less wrapping functions that ensure a proper locking wrt the filesystem and call the generic quota operations.

What filesystem should expect from the generic quota functions:

| ops | FS recursion | Held locks when called |
|---|---|---|
| write_dquot: | yes | dqonoff_sem or dqptr_sem |
| acquire_dquot: | yes | dqonoff_sem or dqptr_sem |
| release_dquot: | yes | dqonoff_sem or dqptr_sem |
| mark_dirty: | no | • |
| write_info: | yes | dqonoff_sem |

FS recursion means calling ->quota_read() and ->quota_write() from superblock operations.

More details about quota locking can be found in fs/dquot.c.

### 1.5.13 vm_operations_struct

prototypes:

```
void (*open)(struct vm_area_struct *);
void (*close)(struct vm_area_struct *);
vm_fault_t (*fault)(struct vm_fault *);
vm_fault_t (*huge_fault)(struct vm_fault *, unsigned int order);
vm_fault_t (*map_pages)(struct vm_fault *, pgoff_t start, pgoff_t end);
vm_fault_t (*page_mkwrite)(struct vm_area_struct *, struct vm_fault *);
vm_fault_t (*pfn_mkwrite)(struct vm_area_struct *, struct vm_fault *);
int (*access)(struct vm_area_struct *, unsigned long, void*, int, int);
```

locking rules:

| ops | mmap_lock | PageLocked(page) |
|---|---|---|
| open: | write | |
| close: | read/write | |
| fault: | read | can return with page locked |
| huge_fault: | maybe-read | |
| map_pages: | maybe-read | |
| page_mkwrite: | read | can return with page locked |
| pfn_mkwrite: | read | |
| access: | read | |

->fault() is called when a previously not present pte is about to be faulted in. The filesystem must find and return the page associated with the passed in "pgoff" in the vm_fault structure. If it is possible that the page may be truncated and/or invalidated, then the filesystem must lock invalidate_lock, then ensure the page is not already truncated (invalidate_lock will block subsequent truncate), and then return with VM_FAULT_LOCKED, and the page locked. The VM will unlock the page.

->huge_fault() is called when there is no PUD or PMD entry present. This gives the filesystem the opportunity to install a PUD or PMD sized page. Filesystems can also use the ->fault method to return a PMD sized page, so implementing this function may not be necessary. In particular, filesystems should not call filemap_fault() from ->huge_fault(). The mmap_lock may not be held when this method is called.

->map_pages() is called when VM asks to map easy accessible pages. Filesystem should find and map pages associated with offsets from "start_pgoff" till "end_pgoff". ->map_pages() is called with the RCU lock held and must not block. If it's not possible to reach a page without blocking, filesystem should skip it. Filesystem should use set_pte_range() to setup page table entry. Pointer to entry associated with the page is passed in "pte" field in vm_fault structure. Pointers to entries for other offsets should be calculated relative to "pte".

->page_mkwrite() is called when a previously read-only pte is about to become writeable. The filesystem again must ensure that there are no truncate/invalidate races or races with operations such as ->remap_file_range or ->copy_file_range, and then return with the page locked.

Usually mapping->invalidate_lock is suitable for proper serialization. If the page has been truncated, the filesystem should not look up a new page like the ->fault() handler, but simply return with VM_FAULT_NOPAGE, which will cause the VM to retry the fault.

->pfn_mkwrite() is the same as page_mkwrite but when the pte is VM_PFNMAP or VM_MIXEDMAP with a page-less entry. Expected return is VM_FAULT_NOPAGE. Or one of the VM_FAULT_ERROR types. The default behavior after this call is to make the pte read-write, unless pfn_mkwrite returns an error.

->access() is called when get_user_pages() fails in access_process_vm(), typically used to debug a process through /proc/pid/mem or ptrace. This function is needed only for VM_IO | VM_PFNMAP VMAs.

---

Dubious stuff

(if you break something or notice that it is broken and do not fix it yourself - at least put it here)

## 1.6 Directory Locking

Locking scheme used for directory operations is based on two kinds of locks - per-inode (->i_rwsem) and per-filesystem (->s_vfs_rename_mutex).

When taking the i_rwsem on multiple non-directory objects, we always acquire the locks in order by increasing address. We'll call that "inode pointer" order in the following.

### 1.6.1 Primitives

For our purposes all operations fall in 6 classes:

1. read access. Locking rules:
   - lock the directory we are accessing (shared)
2. object creation. Locking rules:
   - lock the directory we are accessing (exclusive)
3. object removal. Locking rules:
   - lock the parent (exclusive)
   - find the victim
   - lock the victim (exclusive)
4. link creation. Locking rules:
   - lock the parent (exclusive)
   - check that the source is not a directory
   - lock the source (exclusive; probably could be weakened to shared)
5. rename that is _not_ cross-directory. Locking rules:
   - lock the parent (exclusive)
   - find the source and target

- decide which of the source and target need to be locked. The source needs to be locked if it's a non-directory, target - if it's a non-directory or about to be removed.
- take the locks that need to be taken (exlusive), in inode pointer order if need to take both (that can happen only when both source and target are non-directories - the source because it wouldn't need to be locked otherwise and the target because mixing directory and non-directory is allowed only with RENAME_EXCHANGE, and that won't be removing the target).

6. cross-directory rename. The trickiest in the whole bunch. Locking rules:
   - lock the filesystem
   - if the parents don't have a common ancestor, fail the operation.
   - lock the parents in ”ancestors first” order (exclusive). If neither is an ancestor of the other, lock the parent of source first.
   - find the source and target.
   - verify that the source is not a descendent of the target and target is not a descendent of source; fail the operation otherwise.
   - lock the subdirectories involved (exclusive), source before target.
   - lock the non-directories involved (exclusive), in inode pointer order.

The rules above obviously guarantee that all directories that are going to be read, modified or removed by method will be locked by the caller.

### 1.6.2 Splicing

There is one more thing to consider - splicing. It's not an operation in its own right; it may happen as part of lookup. We speak of the operations on directory trees, but we obviously do not have the full picture of those - especially for network filesystems. What we have is a bunch of subtrees visible in dcache and locking happens on those. Trees grow as we do operations; memory pressure prunes them. Normally that's not a problem, but there is a nasty twist - what should we do when one growing tree reaches the root of another? That can happen in several scenarios, starting from ”somebody mounted two nested subtrees from the same NFS4 server and doing lookups in one of them has reached the root of another”; there's also open-by-fhandle stuff, and there's a possibility that directory we see in one place gets moved by the server to another and we run into it when we do a lookup.

For a lot of reasons we want to have the same directory present in dcache only once. Multiple aliases are not allowed. So when lookup runs into a subdirectory that already has an alias, something needs to be done with dcache trees. Lookup is already holding the parent locked. If alias is a root of separate tree, it gets attached to the directory we are doing a lookup in, under the name we'd been looking for. If the alias is already a child of the directory we are looking in, it changes name to the one we'd been looking for. No extra locking is involved in these two cases. However, if it's a child of some other directory, the things get trickier. First of all, we verify that it is *not* an ancestor of our directory and fail the lookup if it is. Then we try to lock the filesystem and the current parent of the alias. If either trylock fails, we fail the lookup. If trylocks succeed, we detach the alias from its current parent and attach to our directory, under the name we are looking for.

Note that splicing does *not* involve any modification of the filesystem; all we change is the view in dcache. Moreover, holding a directory locked exclusive prevents such changes involving

its children and holding the filesystem lock prevents any changes of tree topology, other than having a root of one tree becoming a child of directory in another. In particular, if two dentries have been found to have a common ancestor after taking the filesystem lock, their relationship will remain unchanged until the lock is dropped. So from the directory operations' point of view splicing is almost irrelevant - the only place where it matters is one step in cross-directory renames; we need to be careful when checking if parents have a common ancestor.

### 1.6.3 Multiple-filesystem stuff

For some filesystems a method can involve a directory operation on another filesystem; it may be ecryptfs doing operation in the underlying filesystem, overlayfs doing something to the layers, network filesystem using a local one as a cache, etc. In all such cases the operations on other filesystems must follow the same locking rules. Moreover, ”a directory operation on this filesystem might involve directory operations on that filesystem” should be an asymmetric relation (or, if you will, it should be possible to rank the filesystems so that directory operation on a filesystem could trigger directory operations only on higher-ranked ones - in these terms overlayfs ranks lower than its layers, network filesystem ranks lower than whatever it caches on, etc.)

### 1.6.4 Deadlock avoidance

If no directory is its own ancestor, the scheme above is deadlock-free.

Proof:

There is a ranking on the locks, such that all primitives take them in order of non-decreasing rank. Namely,

- rank ->i_rwsem of non-directories on given filesystem in inode pointer order.
- put ->i_rwsem of all directories on a filesystem at the same rank, lower than ->i_rwsem of any non-directory on the same filesystem.
- put ->s_vfs_rename_mutex at rank lower than that of any ->i_rwsem on the same filesystem.
- among the locks on different filesystems use the relative rank of those filesystems.

For example, if we have NFS filesystem caching on a local one, we have

1. ->s_vfs_rename_mutex of NFS filesystem
2. ->i_rwsem of directories on that NFS filesystem, same rank for all
3. ->i_rwsem of non-directories on that filesystem, in order of increasing address of inode
4. ->s_vfs_rename_mutex of local filesystem
5. ->i_rwsem of directories on the local filesystem, same rank for all
6. ->i_rwsem of non-directories on local filesystem, in order of increasing address of inode.

It's easy to verify that operations never take a lock with rank lower than that of an already held lock.

Suppose deadlocks are possible. Consider the minimal deadlocked set of threads. It is a cycle of several threads, each blocked on a lock held by the next thread in the cycle.

Since the locking order is consistent with the ranking, all contended locks in the minimal deadlock will be of the same rank, i.e. they all will be ->i_rwsem of directories on the same filesystem. Moreover, without loss of generality we can assume that all operations are done directly to that filesystem and none of them has actually reached the method call.

In other words, we have a cycle of threads, T1,..., Tn, and the same number of directories (D1,...,Dn) such that

T1 is blocked on D1 which is held by T2

T2 is blocked on D2 which is held by T3

...

Tn is blocked on Dn which is held by T1.

Each operation in the minimal cycle must have locked at least one directory and blocked on attempt to lock another. That leaves only 3 possible operations: directory removal (locks parent, then child), same-directory rename killing a subdirectory (ditto) and cross-directory rename of some sort.

There must be a cross-directory rename in the set; indeed, if all operations had been of the ”lock parent, then child” sort we would have Dn a parent of D1, which is a parent of D2, which is a parent of D3, ..., which is a parent of Dn. Relationships couldn't have changed since the moment directory locks had been acquired, so they would all hold simultaneously at the deadlock time and we would have a loop.

Since all operations are on the same filesystem, there can't be more than one cross-directory rename among them. Without loss of generality we can assume that T1 is the one doing a cross-directory rename and everything else is of the ”lock parent, then child” sort.

In other words, we have a cross-directory rename that locked Dn and blocked on attempt to lock D1, which is a parent of D2, which is a parent of D3, ..., which is a parent of Dn. Relationships between D1,...,Dn all hold simultaneously at the deadlock time. Moreover, cross-directory rename does not get to locking any directories until it has acquired filesystem lock and verified that directories involved have a common ancestor, which guarantees that ancestry relationships between all of them had been stable.

Consider the order in which directories are locked by the cross-directory rename; parents first, then possibly their children. Dn and D1 would have to be among those, with Dn locked before D1. Which pair could it be?

It can't be the parents - indeed, since D1 is an ancestor of Dn, it would be the first parent to be locked. Therefore at least one of the children must be involved and thus neither of them could be a descendent of another - otherwise the operation would not have progressed past locking the parents.

It can't be a parent and its child; otherwise we would've had a loop, since the parents are locked before the children, so the parent would have to be a descendent of its child.

It can't be a parent and a child of another parent either. Otherwise the child of the parent in question would've been a descendent of another child.

That leaves only one possibility - namely, both Dn and D1 are among the children, in some order. But that is also impossible, since neither of the children is a descendent of another.

That concludes the proof, since the set of operations with the properties requiered for a minimal deadlock can not exist.

Note that the check for having a common ancestor in cross-directory rename is crucial - without it a deadlock would be possible. Indeed, suppose the parents are initially in different trees; we would lock the parent of source, then try to lock the parent of target, only to have an unrelated lookup splice a distant ancestor of source to some distant descendent of the parent of target. At that point we have cross-directory rename holding the lock on parent of source and trying to lock its distant ancestor. Add a bunch of rmdir() attempts on all directories in between (all of those would fail with -ENOTEMPTY, had they ever gotten the locks) and voila - we have a deadlock.

### 1.6.5 Loop avoidance

These operations are guaranteed to avoid loop creation. Indeed, the only operation that could introduce loops is cross-directory rename. Suppose after the operation there is a loop; since there hadn't been such loops before the operation, at least on of the nodes in that loop must've had its parent changed. In other words, the loop must be passing through the source or, in case of exchange, possibly the target.

Since the operation has succeeded, neither source nor target could have been ancestors of each other. Therefore the chain of ancestors starting in the parent of source could not have passed through the target and vice versa. On the other hand, the chain of ancestors of any node could not have passed through the node itself, or we would've had a loop before the operation. But everything other than source and target has kept the parent after the operation, so the operation does not change the chains of ancestors of (ex-)parents of source and target. In particular, those chains must end after a finite number of steps.

Now consider the loop created by the operation. It passes through either source or target; the next node in the loop would be the ex-parent of target or source resp. After that the loop would follow the chain of ancestors of that parent. But as we have just shown, that chain must end after a finite number of steps, which means that it can't be a part of any loop. Q.E.D.

While this locking scheme works for arbitrary DAGs, it relies on ability to check that directory is a descendent of another object. Current implementation assumes that directory graph is a tree. This assumption is also preserved by all operations (cross-directory rename on a tree that would not introduce a cycle will leave it a tree and link() fails for directories).

Notice that ”directory” in the above == ”anything that might have children”, so if we are going to introduce hybrid objects we will need either to make sure that link(2) doesn't work for them or to make changes in *`is_subdir()`* that would make it work even in presence of such beasts.

## 1.7 The Devpts Filesystem

Each mount of the devpts filesystem is now distinct such that ptys and their indices allocated in one mount are independent from ptys and their indices in all other mounts.

All mounts of the devpts filesystem now create a `/dev/pts/ptmx` node with permissions `0000`.

To retain backwards compatibility the a ptmx device node (aka any node created with `mknod name c 5 2`) when opened will look for an instance of devpts under the name `pts` in the same directory as the ptmx device node.

As an option instead of placing a `/dev/ptmx` device node at `/dev/ptmx` it is possible to place a symlink to `/dev/pts/ptmx` at `/dev/ptmx` or to bind mount `/dev/ptx/ptmx` to `/dev/ptmx`. If

you opt for using the devpts filesystem in this manner devpts should be mounted with the `ptmxmode=0666`, or `chmod 0666 /dev/pts/ptmx` should be called.

Total count of pty pairs in all instances is limited by sysctls:

```
kernel.pty.max = 4096       - global limit
kernel.pty.reserve = 1024   - reserved for filesystems mounted from the␣
↪initial mount namespace
kernel.pty.nr               - current count of ptys
```

Per-instance limit could be set by adding mount option `max=<count>`.

This feature was added in kernel 3.4 together with `sysctl kernel.pty.reserve`.

In kernels older than 3.4 sysctl `kernel.pty.max` works as per-instance limit.

## 1.8 Linux Directory Notification

Stephen Rothwell <sfr@canb.auug.org.au>

The intention of directory notification is to allow user applications to be notified when a directory, or any of the files in it, are changed. The basic mechanism involves the application registering for notification on a directory using a fcntl(2) call and the notifications themselves being delivered using signals.

The application decides which ”events” it wants to be notified about. The currently defined events are:

| | |
|---|---|
| DN_ACCESS | A file in the directory was accessed (read) |
| DN_MODIFY | A file in the directory was modified (write,truncate) |
| DN_CREATE | A file was created in the directory |
| DN_DELETE | A file was unlinked from directory |
| DN_RENAME | A file in the directory was renamed |
| DN_ATTRIB | A file in the directory had its attributes changed (chmod,chown) |

Usually, the application must reregister after each notification, but if DN_MULTISHOT is or'ed with the event mask, then the registration will remain until explicitly removed (by registering for no events).

By default, SIGIO will be delivered to the process and no other useful information. However, if the F_SETSIG fcntl(2) call is used to let the kernel know which signal to deliver, a siginfo structure will be passed to the signal handler and the si_fd member of that structure will contain the file descriptor associated with the directory in which the event occurred.

Preferably the application will choose one of the real time signals (SIGRTMIN + <n>) so that the notifications may be queued. This is especially important if DN_MULTISHOT is specified. Note that SIGRTMIN is often blocked, so it is better to use (at least) SIGRTMIN + 1.

### 1.8.1 Implementation expectations (features and bugs :-))

The notification should work for any local access to files even if the actual file system is on a remote server. This implies that remote access to files served by local user mode servers should be notified. Also, remote accesses to files served by a local kernel NFS server should be notified.

In order to make the impact on the file system code as small as possible, the problem of hard links to files has been ignored. So if a file (x) exists in two directories (a and b) then a change to the file using the name "a/x" should be notified to a program expecting notifications on directory "a", but will not be notified to one expecting notifications on directory "b".

Also, files that are unlinked, will still cause notifications in the last directory that they were linked to.

### 1.8.2 Configuration

Dnotify is controlled via the CONFIG_DNOTIFY configuration option. When disabled, fcntl(fd, F_NOTIFY, ...) will return -EINVAL.

### 1.8.3 Example

See tools/testing/selftests/filesystems/dnotify_test.c for an example.

### 1.8.4 NOTE

Beginning with Linux 2.6.13, dnotify has been replaced by inotify. See *Inotify - A Powerful yet Simple File Change Notification System* for more information on it.

## 1.9 Fiemap Ioctl

The fiemap ioctl is an efficient method for userspace to get file extent mappings. Instead of block-by-block mapping (such as bmap), fiemap returns a list of extents.

### 1.9.1 Request Basics

A fiemap request is encoded within struct fiemap:

```
struct fiemap {
      __u64   fm_start;        /* logical offset (inclusive) at
                                * which to start mapping (in) */
      __u64   fm_length;       /* logical length of mapping which
                                * userspace cares about (in) */
      __u32   fm_flags;        /* FIEMAP_FLAG_* flags for request (in/out) */
      __u32   fm_mapped_extents; /* number of extents that were
                                  * mapped (out) */
      __u32   fm_extent_count; /* size of fm_extents array (in) */
      __u32   fm_reserved;
```

```
      struct fiemap_extent fm_extents[0]; /* array of mapped extents (out) */
};
```

fm_start, and fm_length specify the logical range within the file which the process would like mappings for. Extents returned mirror those on disk - that is, the logical offset of the 1st returned extent may start before fm_start, and the range covered by the last returned extent may end after fm_length. All offsets and lengths are in bytes.

Certain flags to modify the way in which mappings are looked up can be set in fm_flags. If the kernel doesn't understand some particular flags, it will return EBADR and the contents of fm_flags will contain the set of flags which caused the error. If the kernel is compatible with all flags passed, the contents of fm_flags will be unmodified. It is up to userspace to determine whether rejection of a particular flag is fatal to its operation. This scheme is intended to allow the fiemap interface to grow in the future but without losing compatibility with old software.

fm_extent_count specifies the number of elements in the fm_extents[] array that can be used to return extents. If fm_extent_count is zero, then the fm_extents[] array is ignored (no extents will be returned), and the fm_mapped_extents count will hold the number of extents needed in fm_extents[] to hold the file's current mapping. Note that there is nothing to prevent the file from changing between calls to FIEMAP.

The following flags can be set in fm_flags:

**FIEMAP_FLAG_SYNC**
: If this flag is set, the kernel will sync the file before mapping extents.

**FIEMAP_FLAG_XATTR**
: If this flag is set, the extents returned will describe the inodes extended attribute lookup tree, instead of its data tree.

### 1.9.2 Extent Mapping

Extent information is returned within the embedded fm_extents array which userspace must allocate along with the fiemap structure. The number of elements in the fiemap_extents[] array should be passed via fm_extent_count. The number of extents mapped by kernel will be returned via fm_mapped_extents. If the number of fiemap_extents allocated is less than would be required to map the requested range, the maximum number of extents that can be mapped in the fm_extent[] array will be returned and fm_mapped_extents will be equal to fm_extent_count. In that case, the last extent in the array will not complete the requested range and will not have the FIEMAP_EXTENT_LAST flag set (see the next section on extent flags).

Each extent is described by a single fiemap_extent structure as returned in fm_extents:

```
struct fiemap_extent {
        __u64       fe_logical;  /* logical offset in bytes for the start of
                                  * the extent */
        __u64       fe_physical; /* physical offset in bytes for the start
                                  * of the extent */
        __u64       fe_length;   /* length in bytes for the extent */
        __u64       fe_reserved64[2];
        __u32       fe_flags;    /* FIEMAP_EXTENT_* flags for this extent */
        __u32       fe_reserved[3];
};
```

All offsets and lengths are in bytes and mirror those on disk. It is valid for an extents logical offset to start before the request or its logical length to extend past the request. Unless FIEMAP_EXTENT_NOT_ALIGNED is returned, fe_logical, fe_physical, and fe_length will be aligned to the block size of the file system. With the exception of extents flagged as FIEMAP_EXTENT_MERGED, adjacent extents will not be merged.

The fe_flags field contains flags which describe the extent returned. A special flag, FIEMAP_EXTENT_LAST is always set on the last extent in the file so that the process making fiemap calls can determine when no more extents are available, without having to call the ioctl again.

Some flags are intentionally vague and will always be set in the presence of other more specific flags. This way a program looking for a general property does not have to know all existing and future flags which imply that property.

For example, if FIEMAP_EXTENT_DATA_INLINE or FIEMAP_EXTENT_DATA_TAIL are set, FIEMAP_EXTENT_NOT_ALIGNED will also be set. A program looking for inline or tail-packed data can key on the specific flag. Software which simply cares not to try operating on non-aligned extents however, can just key on FIEMAP_EXTENT_NOT_ALIGNED, and not have to worry about all present and future flags which might imply unaligned data. Note that the opposite is not true - it would be valid for FIEMAP_EXTENT_NOT_ALIGNED to appear alone.

**FIEMAP_EXTENT_LAST**
: This is generally the last extent in the file. A mapping attempt past this extent may return nothing. Some implementations set this flag to indicate this extent is the last one in the range queried by the user (via fiemap->fm_length).

**FIEMAP_EXTENT_UNKNOWN**
: The location of this extent is currently unknown. This may indicate the data is stored on an inaccessible volume or that no storage has been allocated for the file yet.

**FIEMAP_EXTENT_DELALLOC**
: This will also set FIEMAP_EXTENT_UNKNOWN.

  Delayed allocation - while there is data for this extent, its physical location has not been allocated yet.

**FIEMAP_EXTENT_ENCODED**
: This extent does not consist of plain filesystem blocks but is encoded (e.g. encrypted or compressed). Reading the data in this extent via I/O to the block device will have undefined results.

Note that it is *always* undefined to try to update the data in-place by writing to the indicated location without the assistance of the filesystem, or to access the data using the information returned by the FIEMAP interface while the filesystem is mounted. In other words, user applications may only read the extent data via I/O to the block device while the filesystem is unmounted, and then only if the FIEMAP_EXTENT_ENCODED flag is clear; user applications must not try reading or writing to the filesystem via the block device under any other circumstances.

**FIEMAP_EXTENT_DATA_ENCRYPTED**
: This will also set FIEMAP_EXTENT_ENCODED The data in this extent has been encrypted by the file system.

**FIEMAP_EXTENT_NOT_ALIGNED**
: Extent offsets and length are not guaranteed to be block aligned.

**FIEMAP_EXTENT_DATA_INLINE**

This will also set FIEMAP_EXTENT_NOT_ALIGNED Data is located within a meta data block.

**FIEMAP_EXTENT_DATA_TAIL**
This will also set FIEMAP_EXTENT_NOT_ALIGNED Data is packed into a block with data from other files.

**FIEMAP_EXTENT_UNWRITTEN**
Unwritten extent - the extent is allocated but its data has not been initialized. This indicates the extent's data will be all zero if read through the filesystem but the contents are undefined if read directly from the device.

**FIEMAP_EXTENT_MERGED**
This will be set when a file does not support extents, i.e., it uses a block based addressing scheme. Since returning an extent for each block back to userspace would be highly inefficient, the kernel will try to merge most adjacent blocks into 'extents'.

### 1.9.3 VFS -> File System Implementation

File systems wishing to support fiemap must implement a ->fiemap callback on their inode_operations structure. The fs ->fiemap call is responsible for defining its set of supported fiemap flags, and calling a helper function on each discovered extent:

```
struct inode_operations {
     ...

     int (*fiemap)(struct inode *, struct fiemap_extent_info *, u64 start,
                   u64 len);
```

->fiemap is passed struct fiemap_extent_info which describes the fiemap request:

```
struct fiemap_extent_info {
      unsigned int fi_flags;          /* Flags as passed from user */
      unsigned int fi_extents_mapped; /* Number of mapped extents */
      unsigned int fi_extents_max;    /* Size of fiemap_extent array */
      struct fiemap_extent *fi_extents_start; /* Start of fiemap_extent array␣
↪*/
};
```

It is intended that the file system should not need to access any of this structure directly. Filesystem handlers should be tolerant to signals and return EINTR once fatal signal received.

Flag checking should be done at the beginning of the ->fiemap callback via the fiemap_prep() helper:

```
int fiemap_prep(struct inode *inode, struct fiemap_extent_info *fieinfo,
                u64 start, u64 *len, u32 supported_flags);
```

The struct fieinfo should be passed in as received from ioctl_fiemap(). The set of fiemap flags which the fs understands should be passed via fs_flags. If fiemap_prep finds invalid user flags, it will place the bad values in fieinfo->fi_flags and return -EBADR. If the file system gets -EBADR, from fiemap_prep(), it should immediately exit, returning that error back to ioctl_fiemap(). Additionally the range is validate against the supported maximum file size.

For each extent in the request range, the file system should call the helper function, fiemap_fill_next_extent():

```
int fiemap_fill_next_extent(struct fiemap_extent_info *info, u64 logical,
                            u64 phys, u64 len, u32 flags, u32 dev);
```

fiemap_fill_next_extent() will use the passed values to populate the next free extent in the fm_extents array. 'General' extent flags will automatically be set from specific flags on behalf of the calling file system so that the userspace API is not broken.

fiemap_fill_next_extent() returns 0 on success, and 1 when the user-supplied fm_extents array is full. If an error is encountered while copying the extent to user memory, -EFAULT will be returned.

## 1.10 File management in the Linux kernel

This document describes how locking for files (struct file) and file descriptor table (struct files) works.

Up until 2.6.12, the file descriptor table has been protected with a lock (files->file_lock) and reference count (files->count). ->file_lock protected accesses to all the file related fields of the table. ->count was used for sharing the file descriptor table between tasks cloned with CLONE_FILES flag. Typically this would be the case for posix threads. As with the common refcounting model in the kernel, the last task doing a put_files_struct() frees the file descriptor (fd) table. The files (struct file) themselves are protected using reference count (->f_count).

In the new lock-free model of file descriptor management, the reference counting is similar, but the locking is based on RCU. The file descriptor table contains multiple elements - the fd sets (open_fds and close_on_exec, the array of file pointers, the sizes of the sets and the array etc.). In order for the updates to appear atomic to a lock-free reader, all the elements of the file descriptor table are in a separate structure - struct fdtable. files_struct contains a pointer to struct fdtable through which the actual fd table is accessed. Initially the fdtable is embedded in files_struct itself. On a subsequent expansion of fdtable, a new fdtable structure is allocated and files->fdtab points to the new structure. The fdtable structure is freed with RCU and lock-free readers either see the old fdtable or the new fdtable making the update appear atomic. Here are the locking rules for the fdtable structure -

1. All references to the fdtable must be done through the files_fdtable() macro:

   ```
   struct fdtable *fdt;

   rcu_read_lock();

   fdt = files_fdtable(files);
   ....
   if (n <= fdt->max_fds)
           ....
   ...
   rcu_read_unlock();
   ```

   files_fdtable() uses rcu_dereference() macro which takes care of the memory barrier requirements for lock-free dereference. The fdtable pointer must be read within the read-

side critical section.

2. Reading of the fdtable as described above must be protected by rcu_read_lock()/rcu_read_unlock().
3. For any update to the fd table, files->file_lock must be held.
4. To look up the file structure given an fd, a reader must use either lookup_fdget_rcu() or files_lookup_fdget_rcu() APIs. These take care of barrier requirements due to lock-free lookup.

   An example:

```
struct file *file;

rcu_read_lock();
file = lookup_fdget_rcu(fd);
rcu_read_unlock();
if (file) {
        ...
        fput(file);
}
....
```

5. Since both fdtable and file structures can be looked up lock-free, they must be installed using rcu_assign_pointer() API. If they are looked up lock-free, rcu_dereference() must be used. However it is advisable to use files_fdtable() and lookup_fdget_rcu()/files_lookup_fdget_rcu() which take care of these issues.
6. While updating, the fdtable pointer must be looked up while holding files->file_lock. If ->file_lock is dropped, then another thread expand the files thereby creating a new fdtable and making the earlier fdtable pointer stale.

   For example:

```
spin_lock(&files->file_lock);
fd = locate_fd(files, file, start);
if (fd >= 0) {
        /* locate_fd() may have expanded fdtable, load the ptr */
        fdt = files_fdtable(files);
        __set_open_fd(fd, fdt);
        __clear_close_on_exec(fd, fdt);
        spin_unlock(&files->file_lock);
.....
```

   Since locate_fd() can drop ->file_lock (and reacquire ->file_lock), the fdtable pointer (fdt) must be loaded after locate_fd().

On newer kernels rcu based file lookup has been switched to rely on SLAB_TYPESAFE_BY_RCU instead of call_rcu(). It isn't sufficient anymore to just acquire a reference to the file in question under rcu using atomic_long_inc_not_zero() since the file might have already been recycled and someone else might have bumped the reference. In other words, callers might see reference count bumps from newer users. For this is reason it is necessary to verify that the pointer is the same before and after the reference count increment. This pattern can be seen in get_file_rcu() and __files_get_rcu().

In addition, it isn't possible to access or check fields in struct file without first aqcuiring a reference on it under rcu lookup. Not doing that was always very dodgy and it was only usable for non-pointer data in struct file. With SLAB_TYPESAFE_BY_RCU it is necessary that callers either first acquire a reference or they must hold the files_lock of the fdtable.

## 1.11 File Locking Release Notes

Andy Walker <andy@lysaker.kvaerner.no>

12 May 1997

### 1.11.1 1. What's New?

#### 1.1 Broken Flock Emulation

The old flock(2) emulation in the kernel was swapped for proper BSD compatible flock(2) support in the 1.3.x series of kernels. With the release of the 2.1.x kernel series, support for the old emulation has been totally removed, so that we don't need to carry this baggage forever.

This should not cause problems for anybody, since everybody using a 2.1.x kernel should have updated their C library to a suitable version anyway (see the file "Documentation/process/changes.rst".)

#### 1.2 Allow Mixed Locks Again

##### 1.2.1 Typical Problems - Sendmail

Because sendmail was unable to use the old flock() emulation, many sendmail installations use fcntl() instead of flock(). This is true of Slackware 3.0 for example. This gave rise to some other subtle problems if sendmail was configured to rebuild the alias file. Sendmail tried to lock the aliases.dir file with fcntl() at the same time as the GDBM routines tried to lock this file with flock(). With pre 1.3.96 kernels this could result in deadlocks that, over time, or under a very heavy mail load, would eventually cause the kernel to lock solid with deadlocked processes.

##### 1.2.2 The Solution

The solution I have chosen, after much experimentation and discussion, is to make flock() and fcntl() locks oblivious to each other. Both can exists, and neither will have any effect on the other.

I wanted the two lock styles to be cooperative, but there were so many race and deadlock conditions that the current solution was the only practical one. It puts us in the same position as, for example, SunOS 4.1.x and several other commercial Unices. The only OS's that support cooperative flock()/fcntl() are those that emulate flock() using fcntl(), with all the problems that implies.

### 1.3 Mandatory Locking As A Mount Option

Mandatory locking was prior to this release a general configuration option that was valid for all mounted filesystems. This had a number of inherent dangers, not the least of which was the ability to freeze an NFS server by asking it to read a file for which a mandatory lock existed.

Such option was dropped in Kernel v5.14.

## 1.12 Filesystem Mount API

### 1.12.1 Overview

The creation of new mounts is now to be done in a multistep process:

(1) Create a filesystem context.

(2) Parse the parameters and attach them to the context. Parameters are expected to be passed individually from userspace, though legacy binary parameters can also be handled.

(3) Validate and pre-process the context.

(4) Get or create a superblock and mountable root.

(5) Perform the mount.

(6) Return an error message attached to the context.

(7) Destroy the context.

To support this, the file_system_type struct gains two new fields:

```
int (*init_fs_context)(struct fs_context *fc);
const struct fs_parameter_description *parameters;
```

The first is invoked to set up the filesystem-specific parts of a filesystem context, including the additional space, and the second points to the parameter description for validation at registration time and querying by a future system call.

Note that security initialisation is done *after* the filesystem is called so that the namespaces may be adjusted first.

### 1.12.2 The Filesystem context

The creation and reconfiguration of a superblock is governed by a filesystem context. This is represented by the fs_context structure:

```
struct fs_context {
        const struct fs_context_operations *ops;
        struct file_system_type *fs_type;
        void                    *fs_private;
        struct dentry           *root;
        struct user_namespace   *user_ns;
        struct net              *net_ns;
        const struct cred       *cred;
```

```
        char                    *source;
        char                    *subtype;
        void                    *security;
        void                    *s_fs_info;
        unsigned int            sb_flags;
        unsigned int            sb_flags_mask;
        unsigned int            s_iflags;
        enum fs_context_purpose purpose:8;
        ...
};
```

The fs_context fields are as follows:

- `const struct fs_context_operations *ops`

  These are operations that can be done on a filesystem context (see below). This must be set by the ->init_fs_context() file_system_type operation.

- `struct file_system_type *fs_type`

  A pointer to the file_system_type of the filesystem that is being constructed or reconfigured. This retains a reference on the type owner.

- `void *fs_private`

  A pointer to the file system's private data. This is where the filesystem will need to store any options it parses.

- `struct dentry *root`

  A pointer to the root of the mountable tree (and indirectly, the superblock thereof). This is filled in by the ->get_tree() op. If this is set, an active reference on root->d_sb must also be held.

- `struct user_namespace *user_ns`
  `struct net *net_ns`

  There are a subset of the namespaces in use by the invoking process. They retain references on each namespace. The subscribed namespaces may be replaced by the filesystem to reflect other sources, such as the parent mount superblock on an automount.

- `const struct cred *cred`

  The mounter's credentials. This retains a reference on the credentials.

- `char *source`

  This specifies the source. It may be a block device (e.g. /dev/sda1) or something more exotic, such as the ”host:/path” that NFS desires.

- `char *subtype`

This is a string to be added to the type displayed in /proc/mounts to qualify it (used by FUSE). This is available for the filesystem to set if desired.

- `void *security`

  A place for the LSMs to hang their security data for the superblock. The relevant security operations are described below.

- `void *s_fs_info`

  The proposed s_fs_info for a new superblock, set in the superblock by *sget_fc()*. This can be used to distinguish superblocks.

- `unsigned int sb_flags`
  `unsigned int sb_flags_mask`

  Which bits SB_* flags are to be set/cleared in super_block::s_flags.

- `unsigned int s_iflags`

  These will be bitwise-OR'd with s->s_iflags when a superblock is created.

- `enum fs_context_purpose`

  This indicates the purpose for which the context is intended. The available values are:

| | |
|---|---|
| FS_CONTEXT_FOR_MOUNT, | New superblock for explicit mount |
| FS_CONTEXT_FOR_SUBMOUNT | New automatic submount of extant mount |
| FS_CONTEXT_FOR_RECONFIGURE | Change an existing mount |

The mount context is created by calling vfs_new_fs_context() or vfs_dup_fs_context() and is destroyed with put_fs_context(). Note that the structure is not refcounted.

VFS, security and filesystem mount options are set individually with vfs_parse_mount_option(). Options provided by the old mount(2) system call as a page of data can be parsed with generic_parse_monolithic().

When mounting, the filesystem is allowed to take data from any of the pointers and attach it to the superblock (or whatever), provided it clears the pointer in the mount context.

The filesystem is also allowed to allocate resources and pin them with the mount context. For instance, NFS might pin the appropriate protocol version module.

### 1.12.3 The Filesystem Context Operations

The filesystem context points to a table of operations:

```
struct fs_context_operations {
        void (*free)(struct fs_context *fc);
        int (*dup)(struct fs_context *fc, struct fs_context *src_fc);
        int (*parse_param)(struct fs_context *fc,
                           struct fs_parameter *param);
        int (*parse_monolithic)(struct fs_context *fc, void *data);
        int (*get_tree)(struct fs_context *fc);
        int (*reconfigure)(struct fs_context *fc);
};
```

These operations are invoked by the various stages of the mount procedure to manage the filesystem context. They are as follows:

- `void (*free)(struct fs_context *fc);`

  Called to clean up the filesystem-specific part of the filesystem context when the context is destroyed. It should be aware that parts of the context may have been removed and NULL'd out by ->get_tree().

- `int (*dup)(struct fs_context *fc, struct fs_context *src_fc);`

  Called when a filesystem context has been duplicated to duplicate the filesystem-private data. An error may be returned to indicate failure to do this.

  > **Warning:** Note that even if this fails, put_fs_context() will be called immediately thereafter, so ->dup() *must* make the filesystem-private data safe for ->free().

- ```
  int (*parse_param)(struct fs_context *fc,
                     struct fs_parameter *param);
  ```

  Called when a parameter is being added to the filesystem context. param points to the key name and maybe a value object. VFS-specific options will have been weeded out and fc->sb_flags updated in the context. Security options will also have been weeded out and fc->security updated.

  The parameter can be parsed with fs_parse() and fs_lookup_param(). Note that the source(s) are presented as parameters named "source".

  If successful, 0 should be returned or a negative error code otherwise.

- `int (*parse_monolithic)(struct fs_context *fc, void *data);`

  Called when the mount(2) system call is invoked to pass the entire data page in one go. If this is expected to be just a list of "key[=val]" items separated by commas, then this may be set to NULL.

  The return value is as for ->parse_param().
  ```

If the filesystem (e.g. NFS) needs to examine the data first and then finds it's the standard key-val list then it may pass it off to generic_parse_monolithic().

- ```
  int (*get_tree)(struct fs_context *fc);
  ```

  Called to get or create the mountable root and superblock, using the information stored in the filesystem context (reconfiguration goes via a different vector). It may detach any resources it desires from the filesystem context and transfer them to the superblock it creates.

  On success it should set fc->root to the mountable root and return 0. In the case of an error, it should return a negative error code.

  The phase on a userspace-driven context will be set to only allow this to be called once on any particular context.

- ```
  int (*reconfigure)(struct fs_context *fc);
  ```

  Called to effect reconfiguration of a superblock using information stored in the filesystem context. It may detach any resources it desires from the filesystem context and transfer them to the superblock. The superblock can be found from fc->root->d_sb.

  On success it should return 0. In the case of an error, it should return a negative error code.

  **Note:** reconfigure is intended as a replacement for remount_fs.

### 1.12.4 Filesystem context Security

The filesystem context contains a security pointer that the LSMs can use for building up a security context for the superblock to be mounted. There are a number of operations used by the new mount code for this purpose:

- ```
  int security_fs_context_alloc(struct fs_context *fc,
                                struct dentry *reference);
  ```

  Called to initialise fc->security (which is preset to NULL) and allocate any resources needed. It should return 0 on success or a negative error code on failure.

  reference will be non-NULL if the context is being created for superblock reconfiguration (FS_CONTEXT_FOR_RECONFIGURE) in which case it indicates the root dentry of the superblock to be reconfigured. It will also be non-NULL in the case of a submount (FS_CONTEXT_FOR_SUBMOUNT) in which case it indicates the automount point.

- ```
  int security_fs_context_dup(struct fs_context *fc,
                              struct fs_context *src_fc);
  ```

  Called to initialise fc->security (which is preset to NULL) and allocate any resources needed. The original filesystem context is pointed to by src_fc and may be used for reference. It should return 0 on success or a negative error code on failure.

- ```
  void security_fs_context_free(struct fs_context *fc);
  ```

  Called to clean up anything attached to fc->security. Note that the contents may have been transferred to a superblock and the pointer cleared during get_tree.

- ```
  int security_fs_context_parse_param(struct fs_context *fc,
                                      struct fs_parameter *param);
  ```

  Called for each mount parameter, including the source. The arguments are as for the ->parse_param() method. It should return 0 to indicate that the parameter should be passed on to the filesystem, 1 to indicate that the parameter should be discarded or an error to indicate that the parameter should be rejected.

  The value pointed to by param may be modified (if a string) or stolen (provided the value pointer is NULL'd out). If it is stolen, 1 must be returned to prevent it being passed to the filesystem.

- ```
  int security_fs_context_validate(struct fs_context *fc);
  ```

  Called after all the options have been parsed to validate the collection as a whole and to do any necessary allocation so that security_sb_get_tree() and security_sb_reconfigure() are less likely to fail. It should return 0 or a negative error code.

  In the case of reconfiguration, the target superblock will be accessible via fc->root.

- ```
  int security_sb_get_tree(struct fs_context *fc);
  ```

  Called during the mount procedure to verify that the specified superblock is allowed to be mounted and to transfer the security data there. It should return 0 or a negative error code.

- ```
  void security_sb_reconfigure(struct fs_context *fc);
  ```

  Called to apply any reconfiguration to an LSM's context. It must not fail. Error checking and resource allocation must be done in advance by the parameter parsing and validation hooks.

- ```
  int security_sb_mountpoint(struct fs_context *fc,
                             struct path *mountpoint,
                             unsigned int mnt_flags);
  ```

  Called during the mount procedure to verify that the root dentry attached to the context is permitted to be attached to the specified mountpoint. It should return 0 on success or a negative error code on failure.

### 1.12.5 VFS Filesystem context API

There are four operations for creating a filesystem context and one for destroying a context:

- ```
  struct fs_context *fs_context_for_mount(struct file_system_type *fs_type,
                                          unsigned int sb_flags);
  ```

  Allocate a filesystem context for the purpose of setting up a new mount, whether that be with a new superblock or sharing an existing one. This sets the superblock flags, initialises the security and calls fs_type->init_fs_context() to initialise the filesystem private data.

  fs_type specifies the filesystem type that will manage the context and sb_flags presets the superblock flags stored therein.

- ```
  struct fs_context *fs_context_for_reconfigure(
          struct dentry *dentry,
          unsigned int sb_flags,
          unsigned int sb_flags_mask);
  ```

  Allocate a filesystem context for the purpose of reconfiguring an existing superblock. dentry provides a reference to the superblock to be configured. sb_flags and sb_flags_mask indicate which superblock flags need changing and to what.

- ```
  struct fs_context *fs_context_for_submount(
          struct file_system_type *fs_type,
          struct dentry *reference);
  ```

  Allocate a filesystem context for the purpose of creating a new mount for an automount point or other derived superblock. fs_type specifies the filesystem type that will manage the context and the reference dentry supplies the parameters. Namespaces are propagated from the reference dentry's superblock also.

  Note that it's not a requirement that the reference dentry be of the same filesystem type as fs_type.

- ```
  struct fs_context *vfs_dup_fs_context(struct fs_context *src_fc);
  ```

  Duplicate a filesystem context, copying any options noted and duplicating or additionally referencing any resources held therein. This is available for use where a filesystem has to get a mount within a mount, such as NFS4 does by internally mounting the root of the target server and then doing a private pathwalk to the target directory.

  The purpose in the new context is inherited from the old one.

- ```
  void put_fs_context(struct fs_context *fc);
  ```

  Destroy a filesystem context, releasing any resources it holds. This calls the ->free() operation. This is intended to be called by anyone who created a filesystem context.

  **Warning:** filesystem contexts are not refcounted, so this causes unconditional destruction.
  ```

In all the above operations, apart from the put op, the return is a mount context pointer or a negative error code.

For the remaining operations, if an error occurs, a negative error code will be returned.

- ```
  int vfs_parse_fs_param(struct fs_context *fc,
                         struct fs_parameter *param);
  ```

  Supply a single mount parameter to the filesystem context. This includes the specification of the source/device which is specified as the "source" parameter (which may be specified multiple times if the filesystem supports that).

  param specifies the parameter key name and the value. The parameter is first checked to see if it corresponds to a standard mount flag (in which case it is used to set an SB_xxx flag and consumed) or a security option (in which case the LSM consumes it) before it is passed on to the filesystem.

  The parameter value is typed and can be one of:

  | | |
  |---|---|
  | fs_value_is_flag | Parameter not given a value |
  | fs_value_is_string | Value is a string |
  | fs_value_is_blob | Value is a binary blob |
  | fs_value_is_filename | Value is a filename* + dirfd |
  | fs_value_is_file | Value is an open file (file*) |

  If there is a value, that value is stored in a union in the struct in one of param->{string,blob,name,file}. Note that the function may steal and clear the pointer, but then becomes responsible for disposing of the object.

- ```
  int vfs_parse_fs_string(struct fs_context *fc, const char *key,
                          const char *value, size_t v_size);
  ```

  A wrapper around vfs_parse_fs_param() that copies the value string it is passed.

- ```
  int generic_parse_monolithic(struct fs_context *fc, void *data);
  ```

  Parse a sys_mount() data page, assuming the form to be a text list consisting of key[=val] options separated by commas. Each item in the list is passed to vfs_mount_option(). This is the default when the ->parse_monolithic() method is NULL.

- ```
  int vfs_get_tree(struct fs_context *fc);
  ```

  Get or create the mountable root and superblock, using the parameters in the filesystem context to select/configure the superblock. This invokes the ->get_tree() method.

- ```
  struct vfsmount *vfs_create_mount(struct fs_context *fc);
  ```

  Create a mount given the parameters in the specified filesystem context. Note that this does not attach the mount to anything.
  ```

### 1.12.6 Superblock Creation Helpers

A number of VFS helpers are available for use by filesystems for the creation or looking up of superblocks.

- ```
  struct super_block *
  sget_fc(struct fs_context *fc,
          int (*test)(struct super_block *sb, struct fs_context *fc),
          int (*set)(struct super_block *sb, struct fs_context *fc));
  ```

  This is the core routine. If test is non-NULL, it searches for an existing superblock matching the criteria held in the fs_context, using the test function to match them. If no match is found, a new superblock is created and the set function is called to set it up.

  Prior to the set function being called, fc->s_fs_info will be transferred to sb->s_fs_info - and fc->s_fs_info will be cleared if set returns success (ie. 0).

The following helpers all wrap *sget_fc()*:

(1) vfs_get_single_super

Only one such superblock may exist in the system. Any further attempt to get a new superblock gets this one (and any parameter differences are ignored).

(2) vfs_get_keyed_super

Multiple superblocks of this type may exist and they're keyed on their s_fs_info pointer (for example this may refer to a namespace).

(3) vfs_get_independent_super

Multiple independent superblocks of this type may exist. This function never matches an existing one and always creates a new one.

### 1.12.7 Parameter Description

Parameters are described using structures defined in linux/fs_parser.h. There's a core description struct that links everything together:

```
struct fs_parameter_description {
        const struct fs_parameter_spec *specs;
        const struct fs_parameter_enum *enums;
};
```

For example:

```
enum {
        Opt_autocell,
        Opt_bar,
        Opt_dyn,
        Opt_foo,
        Opt_source,
};

static const struct fs_parameter_description afs_fs_parameters = {
```
```

```
        .specs          = afs_param_specs,
        .enums          = afs_param_enums,
};
```

The members are as follows:

(1)
```
const struct fs_parameter_specification *specs;
```

Table of parameter specifications, terminated with a null entry, where the entries are of type:

```
struct fs_parameter_spec {
        const char              *name;
        u8                      opt;
        enum fs_parameter_type  type:8;
        unsigned short          flags;
};
```

The 'name' field is a string to match exactly to the parameter key (no wildcards, patterns and no case-independence) and 'opt' is the value that will be returned by the fs_parser() function in the case of a successful match.

The 'type' field indicates the desired value type and must be one of:

| TYPE NAME | EXPECTED VALUE | RESULT IN |
|---|---|---|
| fs_param_is_flag | No value | n/a |
| fs_param_is_bool | Boolean value | result->boolean |
| fs_param_is_u32 | 32-bit unsigned int | result->uint_32 |
| fs_param_is_u32_octal | 32-bit octal int | result->uint_32 |
| fs_param_is_u32_hex | 32-bit hex int | result->uint_32 |
| fs_param_is_s32 | 32-bit signed int | result->int_32 |
| fs_param_is_u64 | 64-bit unsigned int | result->uint_64 |
| fs_param_is_enum | Enum value name | result->uint_32 |
| fs_param_is_string | Arbitrary string | param->string |
| fs_param_is_blob | Binary blob | param->blob |
| fs_param_is_blockdev | Blockdev path | • Needs lookup |
| fs_param_is_path | Path | • Needs lookup |
| fs_param_is_fd | File descriptor | result->int_32 |

Note that if the value is of fs_param_is_bool type, fs_parse() will try to match any string value against ”0”, ”1”, ”no”, ”yes”, ”false”, ”true”.

Each parameter can also be qualified with 'flags':

| | |
|---|---|
| fs_param_v_optional | The value is optional |
| fs_param_neg_with_no | result->negated set if key is prefixed with ”no” |
| fs_param_neg_with_empty | result->negated set if value is ”” |
| fs_param_deprecated | The parameter is deprecated. |

These are wrapped with a number of convenience wrappers:

| MACRO | SPECIFIES |
|---|---|
| fsparam_flag() | fs_param_is_flag |
| fsparam_flag_no() | fs_param_is_flag, fs_param_neg_with_no |
| fsparam_bool() | fs_param_is_bool |
| fsparam_u32() | fs_param_is_u32 |
| fsparam_u32oct() | fs_param_is_u32_octal |
| fsparam_u32hex() | fs_param_is_u32_hex |
| fsparam_s32() | fs_param_is_s32 |
| fsparam_u64() | fs_param_is_u64 |
| fsparam_enum() | fs_param_is_enum |
| fsparam_string() | fs_param_is_string |
| fsparam_blob() | fs_param_is_blob |
| fsparam_bdev() | fs_param_is_blockdev |
| fsparam_path() | fs_param_is_path |
| fsparam_fd() | fs_param_is_fd |

all of which take two arguments, name string and option number - for example:

```
static const struct fs_parameter_spec afs_param_specs[] = {
        fsparam_flag    ("autocell",    Opt_autocell),
        fsparam_flag    ("dyn",         Opt_dyn),
        fsparam_string  ("source",      Opt_source),
        fsparam_flag_no ("foo",         Opt_foo),
        {}
};
```

An addition macro, __fsparam() is provided that takes an additional pair of arguments to specify the type and the flags for anything that doesn't match one of the above macros.

(2)
```
const struct fs_parameter_enum *enums;
```

Table of enum value names to integer mappings, terminated with a null entry. This is of type:

```
struct fs_parameter_enum {
        u8              opt;
        char            name[14];
        u8              value;
};
```

Where the array is an unsorted list of { parameter ID, name }-keyed elements that indicate the value to map to, e.g.:

```
static const struct fs_parameter_enum afs_param_enums[] = {
        { Opt_bar,   "x",      1},
        { Opt_bar,   "y",      23},
        { Opt_bar,   "z",      42},
};
```

If a parameter of type fs_param_is_enum is encountered, fs_parse() will try to look the value up in the enum table and the result will be stored in the parse result.

The parser should be pointed to by the parser pointer in the file_system_type struct as this will provide validation on registration (if CONFIG_VALIDATE_FS_PARSER=y) and will allow the description to be queried from userspace using the fsinfo() syscall.

### 1.12.8 Parameter Helper Functions

A number of helper functions are provided to help a filesystem or an LSM process the parameters it is given.

- ```
  int lookup_constant(const struct constant_table tbl[],
                      const char *name, int not_found);
  ```

  Look up a constant by name in a table of name -> integer mappings. The table is an array of elements of the following type:

  ```
  struct constant_table {
          const char      *name;
          int             value;
  };
  ```

  If a match is found, the corresponding value is returned. If a match isn't found, the not_found value is returned instead.

- ```
  bool validate_constant_table(const struct constant_table *tbl,
                               size_t tbl_size,
                               int low, int high, int special);
  ```

  Validate a constant table. Checks that all the elements are appropriately ordered, that there are no duplicates and that the values are between low and high inclusive, though provision is made for one allowable special value outside of that range. If no special value is required, special should just be set to lie inside the low-to-high range.

  If all is good, true is returned. If the table is invalid, errors are logged to the kernel log buffer and false is returned.

- ```
  bool fs_validate_description(const struct fs_parameter_description *desc);
  ```

  This performs some validation checks on a parameter description. It returns true if the description is good and false if it is not. It will log errors to the kernel log buffer if validation fails.

- ```
  int fs_parse(struct fs_context *fc,
               const struct fs_parameter_description *desc,
               struct fs_parameter *param,
               struct fs_parse_result *result);
  ```

  This is the main interpreter of parameters. It uses the parameter description to look up a parameter by key name and to convert that to an option number (which it returns).

If successful, and if the parameter type indicates the result is a boolean, integer or enum type, the value is converted by this function and the result stored in result->{boolean,int_32,uint_32,uint_64}.

If a match isn't initially made, the key is prefixed with "no" and no value is present then an attempt will be made to look up the key with the prefix removed. If this matches a parameter for which the type has flag fs_param_neg_with_no set, then a match will be made and result->negated will be set to true.

If the parameter isn't matched, -ENOPARAM will be returned; if the parameter is matched, but the value is erroneous, -EINVAL will be returned; otherwise the parameter's option number will be returned.

- ```
  int fs_lookup_param(struct fs_context *fc,
                      struct fs_parameter *value,
                      bool want_bdev,
                      unsigned int flags,
                      struct path *_path);
  ```

  This takes a parameter that carries a string or filename type and attempts to do a path lookup on it. If the parameter expects a blockdev, a check is made that the inode actually represents one.

  Returns 0 if successful and *_path will be set; returns a negative error code if not.

## 1.13 Quota subsystem

Quota subsystem allows system administrator to set limits on used space and number of used inodes (inode is a filesystem structure which is associated with each file or directory) for users and/or groups. For both used space and number of used inodes there are actually two limits. The first one is called softlimit and the second one hardlimit. A user can never exceed a hardlimit for any resource (unless he has CAP_SYS_RESOURCE capability). User is allowed to exceed softlimit but only for limited period of time. This period is called "grace period" or "grace time". When grace time is over, user is not able to allocate more space/inodes until he frees enough of them to get below softlimit.

Quota limits (and amount of grace time) are set independently for each filesystem.

For more details about quota design, see the documentation in quota-tools package (https://sourceforge.net/projects/linuxquota).

### 1.13.1 Quota netlink interface

When user exceeds a softlimit, runs out of grace time or reaches hardlimit, quota subsystem traditionally printed a message to the controlling terminal of the process which caused the excess. This method has the disadvantage that when user is using a graphical desktop he usually cannot see the message. Thus quota netlink interface has been designed to pass information about the above events to userspace. There they can be captured by an application and processed accordingly.

The interface uses generic netlink framework (see https://lwn.net/Articles/208755/ and http://www.infradead.org/~tgr/libnl/ for more details about this layer). The name of the quota generic
  ```

netlink interface is "VFS_DQUOT". Definitions of constants below are in <linux/quota.h>. Since the quota netlink protocol is not namespace aware, quota netlink messages are sent only in initial network namespace.

Currently, the interface supports only one message type QUOTA_NL_C_WARNING. This command is used to send a notification about any of the above mentioned events. Each message has six attributes. These are (type of the argument is in parentheses):

**QUOTA_NL_A_QTYPE (u32)**

- type of quota being exceeded (one of USRQUOTA, GRPQUOTA)

**QUOTA_NL_A_EXCESS_ID (u64)**

- UID/GID (depends on quota type) of user / group whose limit is being exceeded.

**QUOTA_NL_A_CAUSED_ID (u64)**

- UID of a user who caused the event

**QUOTA_NL_A_WARNING (u32)**

- what kind of limit is exceeded:

  **QUOTA_NL_IHARDWARN**
  inode hardlimit

  **QUOTA_NL_ISOFTLONGWARN**
  inode softlimit is exceeded longer than given grace period

  **QUOTA_NL_ISOFTWARN**
  inode softlimit

  **QUOTA_NL_BHARDWARN**
  space (block) hardlimit

  **QUOTA_NL_BSOFTLONGWARN**
  space (block) softlimit is exceeded longer than given grace period.

  **QUOTA_NL_BSOFTWARN**
  space (block) softlimit

- four warnings are also defined for the event when user stops exceeding some limit:

  **QUOTA_NL_IHARDBELOW**
  inode hardlimit

  **QUOTA_NL_ISOFTBELOW**
  inode softlimit

  **QUOTA_NL_BHARDBELOW**
  space (block) hardlimit

  **QUOTA_NL_BSOFTBELOW**
  space (block) softlimit

**QUOTA_NL_A_DEV_MAJOR (u32)**

- major number of a device with the affected filesystem

**QUOTA_NL_A_DEV_MINOR (u32)**

- minor number of a device with the affected filesystem

## 1.14 The seq_file Interface



This file is originally from the LWN.net Driver Porting series at https://lwn.net/Articles/driver-porting/

There are numerous ways for a device driver (or other kernel component) to provide information to the user or system administrator. One useful technique is the creation of virtual files, in debugfs, /proc or elsewhere. Virtual files can provide human-readable output that is easy to get at without any special utility programs; they can also make life easier for script writers. It is not surprising that the use of virtual files has grown over the years.

Creating those files correctly has always been a bit of a challenge, however. It is not that hard to make a virtual file which returns a string. But life gets trickier if the output is long - anything greater than an application is likely to read in a single operation. Handling multiple reads (and seeks) requires careful attention to the reader's position within the virtual file - that position is, likely as not, in the middle of a line of output. The kernel has traditionally had a number of implementations that got this wrong.

The 2.6 kernel contains a set of functions (implemented by Alexander Viro) which are designed to make it easy for virtual file creators to get it right.

The seq_file interface is available via <linux/seq_file.h>. There are three aspects to seq_file:

- An iterator interface which lets a virtual file implementation step through the objects it is presenting.
- Some utility functions for formatting objects for output without needing to worry about things like output buffers.
- A set of canned file_operations which implement most operations on the virtual file.

We'll look at the seq_file interface via an extremely simple example: a loadable module which creates a file called /proc/sequence. The file, when read, simply produces a set of increasing integer values, one per line. The sequence will continue until the user loses patience and finds something better to do. The file is seekable, in that one can do something like the following:

```
dd if=/proc/sequence of=out1 count=1
dd if=/proc/sequence skip=1 of=out2 count=1
```

Then concatenate the output files out1 and out2 and get the right result. Yes, it is a thoroughly useless module, but the point is to show how the mechanism works without getting lost in other details. (Those wanting to see the full source for this module can find it at https://lwn.net/Articles/22359/).

### 1.14.1 Deprecated create_proc_entry

Note that the above article uses create_proc_entry which was removed in kernel 3.10. Current versions require the following update:

```
-   entry = create_proc_entry("sequence", 0, NULL);
-   if (entry)
-           entry->proc_fops = &ct_file_ops;
+   entry = proc_create("sequence", 0, NULL, &ct_file_ops);
```

### 1.14.2 The iterator interface

Modules implementing a virtual file with seq_file must implement an iterator object that allows stepping through the data of interest during a "session" (roughly one read() system call). If the iterator is able to move to a specific position - like the file they implement, though with freedom to map the position number to a sequence location in whatever way is convenient - the iterator need only exist transiently during a session. If the iterator cannot easily find a numerical position but works well with a first/next interface, the iterator can be stored in the private data area and continue from one session to the next.

A seq_file implementation that is formatting firewall rules from a table, for example, could provide a simple iterator that interprets position N as the Nth rule in the chain. A seq_file implementation that presents the content of a, potentially volatile, linked list might record a pointer into that list, providing that can be done without risk of the current location being removed.

Positioning can thus be done in whatever way makes the most sense for the generator of the data, which need not be aware of how a position translates to an offset in the virtual file. The one obvious exception is that a position of zero should indicate the beginning of the file.

The /proc/sequence iterator just uses the count of the next number it will output as its position.

Four functions must be implemented to make the iterator work. The first, called start(), starts a session and takes a position as an argument, returning an iterator which will start reading at that position. The pos passed to start() will always be either zero, or the most recent pos used in the previous session.

For our simple sequence example, the start() function looks like:

```
static void *ct_seq_start(struct seq_file *s, loff_t *pos)
{
        loff_t *spos = kmalloc(sizeof(loff_t), GFP_KERNEL);
        if (! spos)
                return NULL;
        *spos = *pos;
        return spos;
}
```

The entire data structure for this iterator is a single loff_t value holding the current position. There is no upper bound for the sequence iterator, but that will not be the case for most other seq_file implementations; in most cases the start() function should check for a "past end of file" condition and return NULL if need be.

For more complicated applications, the private field of the seq_file structure can be used to hold state from session to session. There is also a special value which can be returned by the start() function called SEQ_START_TOKEN; it can be used if you wish to instruct your show() function (described below) to print a header at the top of the output. SEQ_START_TOKEN should only be used if the offset is zero, however. SEQ_START_TOKEN has no special meaning to the core seq_file code. It is provided as a convenience for a start() function to communicate with the next() and show() functions.

The next function to implement is called, amazingly, next(); its job is to move the iterator forward to the next position in the sequence. The example module can simply increment the position by one; more useful modules will do what is needed to step through some data structure. The next() function returns a new iterator, or NULL if the sequence is complete. Here's the example version:

```
static void *ct_seq_next(struct seq_file *s, void *v, loff_t *pos)
{
        loff_t *spos = v;
        *pos = ++*spos;
        return spos;
}
```

The next() function should set *pos to a value that start() can use to find the new location in the sequence. When the iterator is being stored in the private data area, rather than being reinitialized on each start(), it might seem sufficient to simply set *pos to any non-zero value (zero always tells start() to restart the sequence). This is not sufficient due to historical problems.

Historically, many next() functions have *not* updated *pos at end-of-file. If the value is then used by start() to initialise the iterator, this can result in corner cases where the last entry in the sequence is reported twice in the file. In order to discourage this bug from being resurrected, the core seq_file code now produces a warning if a next() function does not change the value of *pos. Consequently a next() function *must* change the value of *pos, and of course must set it to a non-zero value.

The stop() function closes a session; its job, of course, is to clean up. If dynamic memory is allocated for the iterator, stop() is the place to free it; if a lock was taken by start(), stop() must release that lock. The value that *pos was set to by the last next() call before stop() is remembered, and used for the first start() call of the next session unless lseek() has been called on the file; in that case next start() will be asked to start at position zero:

```
static void ct_seq_stop(struct seq_file *s, void *v)
{
        kfree(v);
}
```

Finally, the show() function should format the object currently pointed to by the iterator for output. The example module's show() function is:

```
static int ct_seq_show(struct seq_file *s, void *v)
{
        loff_t *spos = v;
        seq_printf(s, "%lld\n", (long long)*spos);
        return 0;
}
```

If all is well, the show() function should return zero. A negative error code in the usual manner indicates that something went wrong; it will be passed back to user space. This function can also return SEQ_SKIP, which causes the current item to be skipped; if the show() function has already generated output before returning SEQ_SKIP, that output will be dropped.

We will look at seq_printf() in a moment. But first, the definition of the seq_file iterator is finished by creating a seq_operations structure with the four functions we have just defined:

```
static const struct seq_operations ct_seq_ops = {
        .start = ct_seq_start,
        .next  = ct_seq_next,
        .stop  = ct_seq_stop,
        .show  = ct_seq_show
};
```

This structure will be needed to tie our iterator to the /proc file in a little bit.

It's worth noting that the iterator value returned by start() and manipulated by the other functions is considered to be completely opaque by the seq_file code. It can thus be anything that is useful in stepping through the data to be output. Counters can be useful, but it could also be a direct pointer into an array or linked list. Anything goes, as long as the programmer is aware that things can happen between calls to the iterator function. However, the seq_file code (by design) will not sleep between the calls to start() and stop(), so holding a lock during that time is a reasonable thing to do. The seq_file code will also avoid taking any other locks while the iterator is active.

The iterator value returned by start() or next() is guaranteed to be passed to a subsequent next() or stop() call. This allows resources such as locks that were taken to be reliably released. There is *no* guarantee that the iterator will be passed to show(), though in practice it often will be.

### 1.14.3 Formatted output

The seq_file code manages positioning within the output created by the iterator and getting it into the user's buffer. But, for that to work, that output must be passed to the seq_file code. Some utility functions have been defined which make this task easy.

Most code will simply use seq_printf(), which works pretty much like printk(), but which requires the seq_file pointer as an argument.

For straight character output, the following functions may be used:

```
seq_putc(struct seq_file *m, char c);
seq_puts(struct seq_file *m, const char *s);
seq_escape(struct seq_file *m, const char *s, const char *esc);
```

The first two output a single character and a string, just like one would expect. seq_escape() is like seq_puts(), except that any character in s which is in the string esc will be represented in octal form in the output.

There are also a pair of functions for printing filenames:

```
int seq_path(struct seq_file *m, const struct path *path,
             const char *esc);
```

```
int seq_path_root(struct seq_file *m, const struct path *path,
                  const struct path *root, const char *esc)
```

Here, path indicates the file of interest, and esc is a set of characters which should be escaped in the output. A call to *seq_path()* will output the path relative to the current process's filesystem root. If a different root is desired, it can be used with seq_path_root(). If it turns out that path cannot be reached from root, seq_path_root() returns SEQ_SKIP.

A function producing complicated output may want to check:

```
bool seq_has_overflowed(struct seq_file *m);
```

and avoid further seq_<output> calls if true is returned.

A true return from seq_has_overflowed means that the seq_file buffer will be discarded and the seq_show function will attempt to allocate a larger buffer and retry printing.

### 1.14.4 Making it all work

So far, we have a nice set of functions which can produce output within the seq_file system, but we have not yet turned them into a file that a user can see. Creating a file within the kernel requires, of course, the creation of a set of file_operations which implement the operations on that file. The seq_file interface provides a set of canned operations which do most of the work. The virtual file author still must implement the open() method, however, to hook everything up. The open function is often a single line, as in the example module:

```
static int ct_open(struct inode *inode, struct file *file)
{
        return seq_open(file, &ct_seq_ops);
}
```

Here, the call to *seq_open()* takes the seq_operations structure we created before, and gets set up to iterate through the virtual file.

On a successful open, *seq_open()* stores the struct seq_file pointer in file->private_data. If you have an application where the same iterator can be used for more than one file, you can store an arbitrary pointer in the private field of the seq_file structure; that value can then be retrieved by the iterator functions.

There is also a wrapper function to *seq_open()* called seq_open_private(). It kmallocs a zero filled block of memory and stores a pointer to it in the private field of the seq_file structure, returning 0 on success. The block size is specified in a third parameter to the function, e.g.:

```
static int ct_open(struct inode *inode, struct file *file)
{
        return seq_open_private(file, &ct_seq_ops,
                                sizeof(struct mystruct));
}
```

There is also a variant function, __seq_open_private(), which is functionally identical except that, if successful, it returns the pointer to the allocated memory block, allowing further initialisation e.g.:

```
static int ct_open(struct inode *inode, struct file *file)
{
        struct mystruct *p =
                __seq_open_private(file, &ct_seq_ops, sizeof(*p));

        if (!p)
                return -ENOMEM;

        p->foo = bar; /* initialize my stuff */
                ...
        p->baz = true;

        return 0;
}
```

A corresponding close function, seq_release_private() is available which frees the memory allocated in the corresponding open.

The other operations of interest - read(), llseek(), and release() - are all implemented by the seq_file code itself. So a virtual file's file_operations structure will look like:

```
static const struct file_operations ct_file_ops = {
        .owner   = THIS_MODULE,
        .open    = ct_open,
        .read    = seq_read,
        .llseek  = seq_lseek,
        .release = seq_release
};
```

There is also a seq_release_private() which passes the contents of the seq_file private field to kfree() before releasing the structure.

The final step is the creation of the /proc file itself. In the example code, that is done in the initialization code in the usual way:

```
static int ct_init(void)
{
        struct proc_dir_entry *entry;

        proc_create("sequence", 0, NULL, &ct_file_ops);
        return 0;
}

module_init(ct_init);
```

And that is pretty much it.

### 1.14.5 seq_list

If your file will be iterating through a linked list, you may find these routines useful:

```
struct list_head *seq_list_start(struct list_head *head,
                                 loff_t pos);
struct list_head *seq_list_start_head(struct list_head *head,
                                      loff_t pos);
struct list_head *seq_list_next(void *v, struct list_head *head,
                                loff_t *ppos);
```

These helpers will interpret pos as a position within the list and iterate accordingly. Your start() and next() functions need only invoke the `seq_list_*` helpers with a pointer to the appropriate list_head structure.

### 1.14.6 The extra-simple version

For extremely simple virtual files, there is an even easier interface. A module can define only the show() function, which should create all the output that the virtual file will contain. The file's open() method then calls:

```
int single_open(struct file *file,
                int (*show)(struct seq_file *m, void *p),
                void *data);
```

When output time comes, the show() function will be called once. The data value given to single_open() can be found in the private field of the seq_file structure. When using single_open(), the programmer should use single_release() instead of *seq_release()* in the file_operations structure to avoid a memory leak.

## 1.15 Shared Subtrees

### 1.15.1 1) Overview

Consider the following situation:

A process wants to clone its own namespace, but still wants to access the CD that got mounted recently. Shared subtree semantics provide the necessary mechanism to accomplish the above.

It provides the necessary building blocks for features like per-user-namespace and versioned filesystem.

### 1.15.2 2) Features

Shared subtree provides four different flavors of mounts; struct vfsmount to be precise

a. shared mount

b. slave mount

c. private mount

d. unbindable mount

2a) A shared mount can be replicated to as many mountpoints and all the replicas continue to be exactly same.

Here is an example:

Let's say /mnt has a mount that is shared:

```
mount --make-shared /mnt
```

Note: mount(8) command now supports the --make-shared flag, so the sample 'smount' program is no longer needed and has been removed.

```
# mount --bind /mnt /tmp
```

The above command replicates the mount at /mnt to the mountpoint /tmp and the contents of both the mounts remain identical.

```
#ls /mnt
a b c

#ls /tmp
a b c
```

Now let's say we mount a device at /tmp/a:

```
# mount /dev/sd0  /tmp/a

#ls /tmp/a
t1 t2 t3

#ls /mnt/a
t1 t2 t3
```

Note that the mount has propagated to the mount at /mnt as well.

And the same is true even when /dev/sd0 is mounted on /mnt/a. The contents will be visible under /tmp/a too.

**2b) A slave mount is like a shared mount except that mount and umount events** only propagate towards it.

All slave mounts have a master mount which is a shared.

Here is an example:

Let's say /mnt has a mount which is shared. # mount --make-shared /mnt

Let's bind mount /mnt to /tmp # mount --bind /mnt /tmp

the new mount at /tmp becomes a shared mount and it is a replica of the mount at /mnt.

Now let's make the mount at /tmp; a slave of /mnt # mount --make-slave /tmp

let's mount /dev/sd0 on /mnt/a # mount /dev/sd0 /mnt/a

#ls /mnt/a t1 t2 t3

#ls /tmp/a t1 t2 t3

Note the mount event has propagated to the mount at /tmp

However let's see what happens if we mount something on the mount at /tmp

# mount /dev/sd1 /tmp/b

#ls /tmp/b s1 s2 s3

#ls /mnt/b

Note how the mount event has not propagated to the mount at /mnt

2c) A private mount does not forward or receive propagation.

This is the mount we are familiar with. Its the default type.

2d) A unbindable mount is a unbindable private mount

let's say we have a mount at /mnt and we make it unbindable:

```
   # mount --make-unbindable /mnt

Let's try to bind mount this mount somewhere else::

   # mount --bind /mnt /tmp
   mount: wrong fs type, bad option, bad superblock on /mnt,
          or too many mounted file systems
```

Binding a unbindable mount is a invalid operation.

### 1.15.3 3) Setting mount states

The mount command (util-linux package) can be used to set mount states:

```
mount --make-shared mountpoint
mount --make-slave mountpoint
mount --make-private mountpoint
mount --make-unbindable mountpoint
```

### 1.15.4 4) Use cases

A) A process wants to clone its own namespace, but still wants to access the CD that got mounted recently.

Solution:

The system administrator can make the mount at /cdrom shared:

```
mount --bind /cdrom /cdrom
mount --make-shared /cdrom
```

Now any process that clones off a new namespace will have a mount at /cdrom which is a replica of the same mount in the parent namespace.

So when a CD is inserted and mounted at /cdrom that mount gets propagated to the other mount at /cdrom in all the other clone namespaces.

B) A process wants its mounts invisible to any other process, but still be able to see the other system mounts.

Solution:

To begin with, the administrator can mark the entire mount tree as shareable:

```
mount --make-rshared /
```

A new process can clone off a new namespace. And mark some part of its namespace as slave:

```
mount --make-rslave /myprivatetree
```

Hence forth any mounts within the /myprivatetree done by the process will not show up in any other namespace. However mounts done in the parent namespace under /myprivatetree still shows up in the process's namespace.

Apart from the above semantics this feature provides the building blocks to solve the following problems:

C) Per-user namespace

The above semantics allows a way to share mounts across namespaces. But namespaces are associated with processes. If namespaces are made first class objects with user API to associate/disassociate a namespace with userid, then each user could have his/her own namespace and tailor it to his/her requirements. This needs to be supported in PAM.

D) Versioned files

If the entire mount tree is visible at multiple locations, then an underlying versioning file system can return different versions of the file depending on the path used to access that file.

An example is:

```
mount --make-shared /
mount --rbind / /view/v1
mount --rbind / /view/v2
mount --rbind / /view/v3
mount --rbind / /view/v4
```

and if /usr has a versioning filesystem mounted, then that mount appears at /view/v1/usr, /view/v2/usr, /view/v3/usr and /view/v4/usr too

A user can request v3 version of the file /usr/fs/namespace.c by accessing /view/v3/usr/fs/namespace.c . The underlying versioning filesystem can then decipher that v3 version of the filesystem is being requested and return the corresponding inode.

### 1.15.5 5) Detailed semantics

The section below explains the detailed semantics of bind, rbind, move, mount, umount and clone-namespace operations.

Note: the word 'vfsmount' and the noun 'mount' have been used to mean the same thing, throughout this document.

5a) Mount states

A given mount can be in one of the following states

1) shared
2) slave
3) shared and slave
4) private
5) unbindable

A 'propagation event' is defined as event generated on a vfsmount that leads to mount or unmount actions in other vfsmounts.

A 'peer group' is defined as a group of vfsmounts that propagate events to each other.

(1) Shared mounts

A 'shared mount' is defined as a vfsmount that belongs to a 'peer group'.

For example:

```
mount --make-shared /mnt
mount --bind /mnt /tmp
```

The mount at /mnt and that at /tmp are both shared and belong to the same peer group. Anything mounted or unmounted under /mnt or /tmp reflect in all the other mounts of its peer group.

(2) Slave mounts

A 'slave mount' is defined as a vfsmount that receives propagation events and does not forward propagation events.

A slave mount as the name implies has a master mount from which mount/unmount events are received. Events do not propagate from the slave mount to the master. Only a shared mount can be made a slave by executing the following command:

```
mount --make-slave mount
```

A shared mount that is made as a slave is no more shared unless modified to become shared.

(3) Shared and Slave

A vfsmount can be both shared as well as slave. This state indicates that the mount is a slave of some vfsmount, and has its own peer group too. This vfsmount receives propagation events from its master vfsmount, and also forwards propagation events to its 'peer group' and to its slave vfsmounts.

Strictly speaking, the vfsmount is shared having its own peer group, and this peer-group is a slave of some other peer group.

Only a slave vfsmount can be made as 'shared and slave' by either executing the following command:

```
mount --make-shared mount
```

or by moving the slave vfsmount under a shared vfsmount.

(4) Private mount

A 'private mount' is defined as vfsmount that does not receive or forward any propagation events.

(5) Unbindable mount

A 'unbindable mount' is defined as vfsmount that does not receive or forward any propagation events and cannot be bind mounted.

State diagram:

The state diagram below explains the state transition of a mount, in response to various commands:

```
------------------------------------------------------------------------
|             |make-shared |  make-slave  | make-private |make-
↪unbindab|
--------------|------------|--------------|--------------|-------------
↪|
|shared       |shared      |*slave/private|   private    | unbindable ␣
↪|
|             |            |              |              |            ␣
↪|
|-------------|------------|--------------|--------------|-------------
↪|
|slave        |shared      | **slave      |    private   | unbindable ␣
↪|
|             |and slave   |              |              |            ␣
↪|
```

```
|-------------|------------|--------------|--------------|-------------
 ↪|
|shared       |shared      | slave        |    private   | unbindable ␣
 ↪|
|and slave    |and slave   |              |              |           ␣
 ↪|
|-------------|------------|--------------|--------------|-------------
 ↪|
|private      |shared      |  **private   |    private   | unbindable ␣
 ↪|
|-------------|------------|--------------|--------------|-------------
 ↪|
|unbindable   |shared      |**unbindable  |    private   | unbindable ␣
 ↪|
-------------------------------------------------------------------------
 ↪-

* if the shared mount is the only mount in its peer group, making it
slave, makes it private automatically. Note that there is no master to
which it can be slaved to.

** slaving a non-shared mount has no effect on the mount.
```

Apart from the commands listed below, the 'move' operation also changes the state of a mount depending on type of the destination mount. Its explained in section 5d.

5b) Bind semantics

Consider the following command:

```
mount --bind A/a  B/b
```

where 'A' is the source mount, 'a' is the dentry in the mount 'A', 'B' is the destination mount and 'b' is the dentry in the destination mount.

The outcome depends on the type of mount of 'A' and 'B'. The table below contains quick reference:

```
-------------------------------------------------------------------
 ↪----------
|         BIND MOUNT OPERATION                                    ␣
 ↪         |
|*****************************************************************|
|source(A)->| shared      |       private  |       slave    |␣
 ↪unbindable |
| dest(B)  |              |                |                |  ␣
 ↪         |
|   |      |              |                |                |  ␣
 ↪         |
|   v      |              |                |                |  ␣
 ↪         |
|*****************************************************************|
```

```
|  shared  | shared       |     shared     | shared & slave | ␣
↪invalid   |
|          |              |                |                |  ␣
↪          |
|non-shared| shared       |      private   |      slave     | ␣
↪invalid   |
*****************************************************************************
```

Details:

1. **'A' is a shared mount and 'B' is a shared mount. A new mount 'C'** which is clone of 'A', is created. Its root dentry is 'a' . 'C' is mounted on mount 'B' at dentry 'b'. Also new mount 'C1', 'C2', 'C3' ... are created and mounted at the dentry 'b' on all mounts where 'B' propagates to. A new propagation tree containing 'C1',..,'Cn' is created. This propagation tree is identical to the propagation tree of 'B'. And finally the peer-group of 'C' is merged with the peer group of 'A'.

2. **'A' is a private mount and 'B' is a shared mount. A new mount 'C'** which is clone of 'A', is created. Its root dentry is 'a'. 'C' is mounted on mount 'B' at dentry 'b'. Also new mount 'C1', 'C2', 'C3' ... are created and mounted at the dentry 'b' on all mounts where 'B' propagates to. A new propagation tree is set containing all new mounts 'C', 'C1', .., 'Cn' with exactly the same configuration as the propagation tree for 'B'.

3. **'A' is a slave mount of mount 'Z' and 'B' is a shared mount. A new** mount 'C' which is clone of 'A', is created. Its root dentry is 'a' . 'C' is mounted on mount 'B' at dentry 'b'. Also new mounts 'C1', 'C2', 'C3' ... are created and mounted at the dentry 'b' on all mounts where 'B' propagates to. A new propagation tree containing the new mounts 'C','C1',.. 'Cn' is created. This propagation tree is identical to the propagation tree for 'B'. And finally the mount 'C' and its peer group is made the slave of mount 'Z'. In other words, mount 'C' is in the state 'slave and shared'.

4. **'A' is a unbindable mount and 'B' is a shared mount. This is a** invalid operation.

5. **'A' is a private mount and 'B' is a non-shared(private or slave or** unbindable) mount. A new mount 'C' which is clone of 'A', is created. Its root dentry is 'a'. 'C' is mounted on mount 'B' at dentry 'b'.

6. **'A' is a shared mount and 'B' is a non-shared mount. A new mount 'C'** which is a clone of 'A' is created. Its root dentry is 'a'. 'C' is mounted on mount 'B' at dentry 'b'. 'C' is made a member of the peer-group of 'A'.

7. **'A' is a slave mount of mount 'Z' and 'B' is a non-shared mount. A** new mount 'C' which is a clone of 'A' is created. Its root dentry is 'a'. 'C' is mounted on mount 'B' at dentry 'b'. Also 'C' is set as a slave mount of 'Z'. In other words 'A' and 'C' are both slave mounts of 'Z'. All mount/unmount events on 'Z' propagates to 'A' and 'C'. But mount/unmount on 'A' do not propagate anywhere else. Similarly mount/unmount on 'C' do not propagate anywhere else.

8. **'A' is a unbindable mount and 'B' is a non-shared mount. This is a** invalid operation. A unbindable mount cannot be bind mounted.

5c) Rbind semantics

rbind is same as bind. Bind replicates the specified mount. Rbind replicates all the mounts in the tree belonging to the specified mount. Rbind mount is bind mount applied to all the mounts in the tree.

If the source tree that is rbind has some unbindable mounts, then the subtree under the unbindable mount is pruned in the new location.

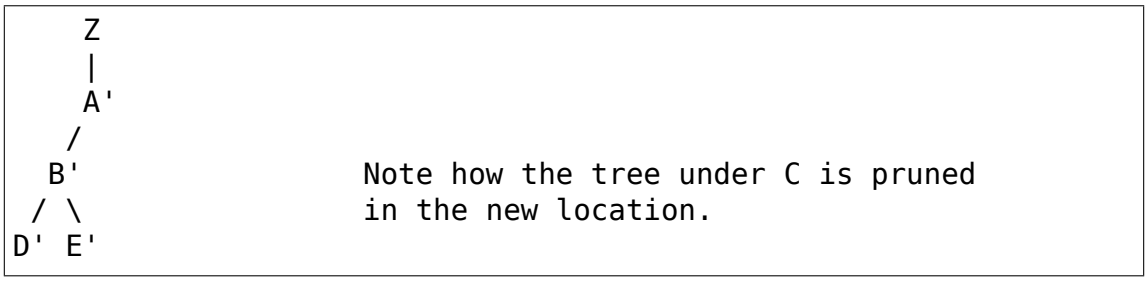
eg:

let's say we have the following mount tree:


Let's say all the mount except the mount C in the tree are of a type other than unbindable.

If this tree is rbound to say Z

We will have the following tree at the new location:




5d) Move semantics

Consider the following command

mount --move A B/b

where 'A' is the source mount, 'B' is the destination mount and 'b' is the dentry in the destination mount.

The outcome depends on the type of the mount of 'A' and 'B'. The table below is a quick reference:

| MOVE MOUNT OPERATION | | | |
|---|---|---|---|
| source(A)->| shared | private | slave | unbindable |
| dest(B) | | | |
| | | | |
| | | | |

| v | | | |
|---|---|---|---|
| shared | shared | shared | shared and slave |
| non-shared | shared | private | slave |
| | unbindable | | |

**Note:** moving a mount residing under a shared mount is invalid.

Details follow:

1. **'A' is a shared mount and 'B' is a shared mount. The mount 'A' is** mounted on mount 'B' at dentry 'b'. Also new mounts 'A1', 'A2'...'An' are created and mounted at dentry 'b' on all mounts that receive propagation from mount 'B'. A new propagation tree is created in the exact same configuration as that of 'B'. This new propagation tree contains all the new mounts 'A1', 'A2'... 'An'. And this new propagation tree is appended to the already existing propagation tree of 'A'.

2. **'A' is a private mount and 'B' is a shared mount. The mount 'A' is** mounted on mount 'B' at dentry 'b'. Also new mount 'A1', 'A2'... 'An' are created and mounted at dentry 'b' on all mounts that receive propagation from mount 'B'. The mount 'A' becomes a shared mount and a propagation tree is created which is identical to that of 'B'. This new propagation tree contains all the new mounts 'A1', 'A2'... 'An'.

3. **'A' is a slave mount of mount 'Z' and 'B' is a shared mount. The** mount 'A' is mounted on mount 'B' at dentry 'b'. Also new mounts 'A1', 'A2'... 'An' are created and mounted at dentry 'b' on all mounts that receive propagation from mount 'B'. A new propagation tree is created in the exact same configuration as that of 'B'. This new propagation tree contains all the new mounts 'A1', 'A2'... 'An'. And this new propagation tree is appended to the already existing propagation tree of 'A'. Mount 'A' continues to be the slave mount of 'Z' but it also becomes 'shared'.

4. **'A' is a unbindable mount and 'B' is a shared mount. The operation** is invalid. Because mounting anything on the shared mount 'B' can create new mounts that get mounted on the mounts that receive propagation from 'B'. And since the mount 'A' is unbindable, cloning it to mount at other mount-points is not possible.

5. **'A' is a private mount and 'B' is a non-shared(private or slave or** unbindable) mount. The mount 'A' is mounted on mount 'B' at dentry 'b'.

6. **'A' is a shared mount and 'B' is a non-shared mount. The mount 'A'** is mounted on mount 'B' at dentry 'b'. Mount 'A' continues to be a shared mount.

7. **'A' is a slave mount of mount 'Z' and 'B' is a non-shared mount.**

The mount 'A' is mounted on mount 'B' at dentry 'b'. Mount 'A' continues to be a slave mount of mount 'Z'.

8. **'A' is a unbindable mount and 'B' is a non-shared mount. The mount** 'A' is mounted on mount 'B' at dentry 'b'. Mount 'A' continues to be a unbindable mount.

5e) Mount semantics

Consider the following command:

```
mount device  B/b
```

'B' is the destination mount and 'b' is the dentry in the destination mount.

The above operation is the same as bind operation with the exception that the source mount is always a private mount.

5f) Unmount semantics

Consider the following command:

```
umount A
```

where 'A' is a mount mounted on mount 'B' at dentry 'b'.

If mount 'B' is shared, then all most-recently-mounted mounts at dentry 'b' on mounts that receive propagation from mount 'B' and does not have sub-mounts within them are unmounted.

Example: Let's say 'B1', 'B2', 'B3' are shared mounts that propagate to each other.

let's say 'A1', 'A2', 'A3' are first mounted at dentry 'b' on mount 'B1', 'B2' and 'B3' respectively.

let's say 'C1', 'C2', 'C3' are next mounted at the same dentry 'b' on mount 'B1', 'B2' and 'B3' respectively.

if 'C1' is unmounted, all the mounts that are most-recently-mounted on 'B1' and on the mounts that 'B1' propagates-to are unmounted.

'B1' propagates to 'B2' and 'B3'. And the most recently mounted mount on 'B2' at dentry 'b' is 'C2', and that of mount 'B3' is 'C3'.

So all 'C1', 'C2' and 'C3' should be unmounted.

If any of 'C2' or 'C3' has some child mounts, then that mount is not unmounted, but all other mounts are unmounted. However if 'C1' is told to be unmounted and 'C1' has some sub-mounts, the umount operation is failed entirely.

5g) Clone Namespace

A cloned namespace contains all the mounts as that of the parent namespace.

Let's say 'A' and 'B' are the corresponding mounts in the parent and the child namespace.

If 'A' is shared, then 'B' is also shared and 'A' and 'B' propagate to each other.

If 'A' is a slave mount of 'Z', then 'B' is also the slave mount of 'Z'.

If 'A' is a private mount, then 'B' is a private mount too.

If 'A' is unbindable mount, then 'B' is a unbindable mount too.

### 1.15.6 6) Quiz

A. What is the result of the following command sequence?

```
mount --bind /mnt /mnt
mount --make-shared /mnt
mount --bind /mnt /tmp
mount --move /tmp /mnt/1
```

what should be the contents of /mnt /mnt/1 /mnt/1/1 should be? Should they all be identical? or should /mnt and /mnt/1 be identical only?

B. What is the result of the following command sequence?

```
mount --make-rshared /
mkdir -p /v/1
mount --rbind / /v/1
```

what should be the content of /v/1/v/1 be?

C. What is the result of the following command sequence?

```
mount --bind /mnt /mnt
mount --make-shared /mnt
mkdir -p /mnt/1/2/3 /mnt/1/test
mount --bind /mnt/1 /tmp
mount --make-slave /mnt
mount --make-shared /mnt
mount --bind /mnt/1/2 /tmp1
mount --make-slave /mnt
```

At this point we have the first mount at /tmp and its root dentry is 1. Let's call this mount 'A' And then we have a second mount at /tmp1 with root dentry 2. Let's call this mount 'B' Next we have a third mount at /mnt with root dentry mnt. Let's call this mount 'C'

'B' is the slave of 'A' and 'C' is a slave of 'B' A -> B -> C

at this point if we execute the following command

mount --bind /bin /tmp/test

The mount is attempted on 'A'

will the mount propagate to 'B' and 'C' ?

what would be the contents of /mnt/1/test be?

### 1.15.7 7) FAQ

**Q1. Why is bind mount needed? How is it different from symbolic links?**
symbolic links can get stale if the destination mount gets unmounted or moved. Bind mounts continue to exist even if the other mount is unmounted or moved.

Q2. Why can't the shared subtree be implemented using exportfs?

exportfs is a heavyweight way of accomplishing part of what shared subtree can do. I cannot imagine a way to implement the semantics of slave mount using exportfs?

Q3 Why is unbindable mount needed?

Let's say we want to replicate the mount tree at multiple locations within the same subtree.

if one rbind mounts a tree within the same subtree 'n' times the number of mounts created is an exponential function of 'n'. Having unbindable mount can help prune the unneeded bind mounts. Here is an example.

**step 1:**
let's say the root tree has just two directories with one vfsmount:

```
                root
               /    \
              tmp    usr

And we want to replicate the tree at multiple
mountpoints under /root/tmp
```

**step 2:**

```
mount --make-shared /root

mkdir -p /tmp/m1

mount --rbind /root /tmp/m1
```

the new tree now looks like this:

```
          root
         /    \
       tmp    usr
      /
     m1
    /  \
   tmp  usr
   /
  m1

it has two vfsmounts
```

**step 3:**

```
    mkdir -p /tmp/m2
    mount --rbind /root /tmp/m2

 the new tree now looks like this::

                root
               /    \
             tmp     usr
            /    \
          m1       m2
         / \       /  \
       tmp  usr   tmp  usr
       / \          /
      m1  m2       m1
          / \      /  \
        tmp usr  tmp   usr
        /        / \
       m1       m1  m2
      /  \
    tmp   usr
    /  \
  m1   m2

it has 6 vfsmounts
```

**step 4:**

**::**

mkdir -p /tmp/m3 mount --rbind /root /tmp/m3

I won't draw the tree..but it has 24 vfsmounts

at step i the number of vfsmounts is V[i] = i*V[i-1]. This is an exponential function. And this tree has way more mounts than what we really needed in the first place.

One could use a series of umount at each step to prune out the unneeded mounts. But there is a better solution. Unclonable mounts come in handy here.

**step 1:**

let's say the root tree has just two directories with one vfsmount:

```
                root
               /    \
              tmp    usr

How do we set up the same tree at multiple locations under
/root/tmp
```

**step 2:**

```
mount --bind /root/tmp /root/tmp

mount --make-rshared /root
mount --make-unbindable /root/tmp

mkdir -p /tmp/m1

mount --rbind /root /tmp/m1
```

the new tree now looks like this:

```
       root
      /    \
    tmp    usr
   /
  m1
 /  \
tmp  usr
```

**step 3:**

```
mkdir -p /tmp/m2
mount --rbind /root /tmp/m2
```

the new tree now looks like this:

```
       root
      /    \
    tmp    usr
   /   \
  m1     m2
 /  \     / \
tmp  usr tmp usr
```

**step 4:**

```
mkdir -p /tmp/m3
mount --rbind /root /tmp/m3
```

the new tree now looks like this:

```
             root
         /           \
        tmp           usr
    /    \    \
  m1     m2     m3
 /  \     / \    /  \
tmp  usr tmp usr tmp usr
```

### 1.15.8 8) Implementation

8A) Datastructure

4 new fields are introduced to struct vfsmount:

- ->mnt_share
- ->mnt_slave_list
- ->mnt_slave
- ->mnt_master

**->mnt_share**
links together all the mount to/from which this vfsmount send/receives propagation events.

**->mnt_slave_list**
links all the mounts to which this vfsmount propagates to.

**->mnt_slave**
links together all the slaves that its master vfsmount propagates to.

**->mnt_master**
points to the master vfsmount from which this vfsmount receives propagation.

**->mnt_flags**
takes two more flags to indicate the propagation status of the vfsmount. MNT_SHARE indicates that the vfsmount is a shared vfsmount. MNT_UNCLONABLE indicates that the vfsmount cannot be replicated.

All the shared vfsmounts in a peer group form a cyclic list through ->mnt_share.

All vfsmounts with the same ->mnt_master form on a cyclic list anchored in ->mnt_master->mnt_slave_list and going through ->mnt_slave.

->mnt_master can point to arbitrary (and possibly different) members of master peer group. To find all immediate slaves of a peer group you need to go through _all_ ->mnt_slave_list of its members. Conceptually it's just a single set - distribution among the individual lists does not affect propagation or the way propagation tree is modified by operations.

All vfsmounts in a peer group have the same ->mnt_master. If it is non-NULL, they form a contiguous (ordered) segment of slave list.

A example propagation tree looks as shown in the figure below. [ NOTE: Though it looks like a forest, if we consider all the shared mounts as a conceptual entity called 'pnode', it becomes a tree]:

In the above figure A,B,C and D all are shared and propagate to each other. 'A' has got 3 slave mounts 'E' 'F' and 'G' 'C' has got 2 slave mounts 'J' and 'K' and 'D' has got two slave mounts 'H' and 'I'. 'E' is also shared with 'K' and they propagate to each other. And 'K' has 3 slaves 'M', 'L' and 'N'

A's ->mnt_share links with the ->mnt_share of 'B' 'C' and 'D'

A's ->mnt_slave_list links with ->mnt_slave of 'E', 'K', 'F' and 'G'

E's ->mnt_share links with ->mnt_share of K

'E', 'K', 'F', 'G' have their ->mnt_master point to struct vfsmount of 'A'

'M', 'L', 'N' have their ->mnt_master point to struct vfsmount of 'K'

K's ->mnt_slave_list links with ->mnt_slave of 'M', 'L' and 'N'

C's ->mnt_slave_list links with ->mnt_slave of 'J' and 'K'

J and K's ->mnt_master points to struct vfsmount of C

and finally D's ->mnt_slave_list links with ->mnt_slave of 'H' and 'I'

'H' and 'I' have their ->mnt_master pointing to struct vfsmount of 'D'.

NOTE: The propagation tree is orthogonal to the mount tree.

8B Locking:

->mnt_share, ->mnt_slave, ->mnt_slave_list, ->mnt_master are protected by namespace_sem (exclusive for modifications, shared for reading).

Normally we have ->mnt_flags modifications serialized by vfsmount_lock. There are two exceptions: do_add_mount() and clone_mnt(). The former modifies a vfsmount that has not been visible in any shared data structures yet. The latter holds namespace_sem and the only references to vfsmount are in lists that can't be traversed without namespace_sem.

8C Algorithm:

The crux of the implementation resides in rbind/move operation.

The overall algorithm breaks the operation into 3 phases: (look at attach_recursive_mnt() and propagate_mnt())

1. prepare phase.
2. commit phases.
3. abort phases.

Prepare phase:

for each mount in the source tree:

a) **Create the necessary number of mount trees to** be attached to each of the mounts that receive propagation from the destination mount.

b) Do not attach any of the trees to its destination. However note down its ->mnt_parent and ->mnt_mountpoint

c) Link all the new mounts to form a propagation tree that is identical to the propagation tree of the destination mount.

If this phase is successful, there should be 'n' new propagation trees; where 'n' is the number of mounts in the source tree. Go to the commit phase

Also there should be 'm' new mount trees, where 'm' is the number of mounts to which the destination mount propagates to.

if any memory allocations fail, go to the abort phase.

**Commit phase**
attach each of the mount trees to their corresponding destination mounts.

**Abort phase**
delete all the newly created trees.

---

**Note:** all the propagation related functionality resides in the file pnode.c

---

version 0.1 (created the initial document, Ram Pai linuxram@us.ibm.com)

version 0.2 (Incorporated comments from Al Viro)

## 1.16 Idmappings

Most filesystem developers will have encountered idmappings. They are used when reading from or writing ownership to disk, reporting ownership to userspace, or for permission checking. This document is aimed at filesystem developers that want to know how idmappings work.

### 1.16.1 Formal notes

An idmapping is essentially a translation of a range of ids into another or the same range of ids. The notational convention for idmappings that is widely used in userspace is:

```
u:k:r
```

`u` indicates the first element in the upper idmapset `U` and `k` indicates the first element in the lower idmapset `K`. The `r` parameter indicates the range of the idmapping, i.e. how many ids are mapped. From now on, we will always prefix ids with `u` or `k` to make it clear whether we're talking about an id in the upper or lower idmapset.

To see what this looks like in practice, let's take the following idmapping:

```
u22:k10000:r3
```

and write down the mappings it will generate:

```
u22 -> k10000
u23 -> k10001
u24 -> k10002
```

From a mathematical viewpoint U and K are well-ordered sets and an idmapping is an order isomorphism from U into K. So U and K are order isomorphic. In fact, U and K are always well-ordered subsets of the set of all possible ids usable on a given system.

Looking at this mathematically briefly will help us highlight some properties that make it easier to understand how we can translate between idmappings. For example, we know that the inverse idmapping is an order isomorphism as well:

```
k10000 -> u22
k10001 -> u23
k10002 -> u24
```

Given that we are dealing with order isomorphisms plus the fact that we're dealing with subsets we can embed idmappings into each other, i.e. we can sensibly translate between different idmappings. For example, assume we've been given the three idmappings:

```
1. u0:k10000:r10000
2. u0:k20000:r10000
3. u0:k30000:r10000
```

and id k11000 which has been generated by the first idmapping by mapping u1000 from the upper idmapset down to k11000 in the lower idmapset.

Because we're dealing with order isomorphic subsets it is meaningful to ask what id k11000 corresponds to in the second or third idmapping. The straightforward algorithm to use is to apply the inverse of the first idmapping, mapping k11000 up to u1000. Afterwards, we can map u1000 down using either the second idmapping mapping or third idmapping mapping. The second idmapping would map u1000 down to 21000. The third idmapping would map u1000 down to u31000.

If we were given the same task for the following three idmappings:

```
1. u0:k10000:r10000
2. u0:k20000:r200
3. u0:k30000:r300
```

we would fail to translate as the sets aren't order isomorphic over the full range of the first idmapping anymore (However they are order isomorphic over the full range of the second idmapping.). Neither the second or third idmapping contain u1000 in the upper idmapset U. This is equivalent to not having an id mapped. We can simply say that u1000 is unmapped in the second and third idmapping. The kernel will report unmapped ids as the overflowuid (uid_t)-1 or overflowgid (gid_t)-1 to userspace.

The algorithm to calculate what a given id maps to is pretty simple. First, we need to verify that the range can contain our target id. We will skip this step for simplicity. After that if we want to know what id maps to we can do simple calculations:

- If we want to map from left to right:

  ```
  u:k:r
  id - u + k = n
  ```

- If we want to map from right to left:

```
u:k:r
id - k + u = n
```

Instead of ”left to right” we can also say ”down” and instead of ”right to left” we can also say ”up”. Obviously mapping down and up invert each other.

To see whether the simple formulas above work, consider the following two idmappings:

```
1. u0:k20000:r10000
2. u500:k30000:r10000
```

Assume we are given `k21000` in the lower idmapset of the first idmapping. We want to know what id this was mapped from in the upper idmapset of the first idmapping. So we're mapping up in the first idmapping:

```
id     - k      + u  = n
k21000 - k20000 + u0 = u1000
```

Now assume we are given the id `u1100` in the upper idmapset of the second idmapping and we want to know what this id maps down to in the lower idmapset of the second idmapping. This means we're mapping down in the second idmapping:

```
id    - u    + k      = n
u1100 - u500 + k30000 = k30600
```

### 1.16.2 General notes

In the context of the kernel an idmapping can be interpreted as mapping a range of userspace ids into a range of kernel ids:

```
userspace-id:kernel-id:range
```

A userspace id is always an element in the upper idmapset of an idmapping of type `uid_t` or `gid_t` and a kernel id is always an element in the lower idmapset of an idmapping of type `kuid_t` or `kgid_t`. From now on ”userspace id” will be used to refer to the well known `uid_t` and `gid_t` types and ”kernel id” will be used to refer to `kuid_t` and `kgid_t`.

The kernel is mostly concerned with kernel ids. They are used when performing permission checks and are stored in an inode's `i_uid` and `i_gid` field. A userspace id on the other hand is an id that is reported to userspace by the kernel, or is passed by userspace to the kernel, or a raw device id that is written or read from disk.

Note that we are only concerned with idmappings as the kernel stores them not how userspace would specify them.

For the rest of this document we will prefix all userspace ids with `u` and all kernel ids with `k`. Ranges of idmappings will be prefixed with `r`. So an idmapping will be written as `u0:k10000:r10000`.

For example, within this idmapping, the id `u1000` is an id in the upper idmapset or ”userspace idmapset” starting with `u0`. And it is mapped to `k11000` which is a kernel id in the lower idmapset or ”kernel idmapset” starting with `k10000`.

A kernel id is always created by an idmapping. Such idmappings are associated with user namespaces. Since we mainly care about how idmappings work we're not going to be concerned with how idmappings are created nor how they are used outside of the filesystem context. This is best left to an explanation of user namespaces.

The initial user namespace is special. It always has an idmapping of the following form:

```
u0:k0:r4294967295
```

which is an identity idmapping over the full range of ids available on this system.

Other user namespaces usually have non-identity idmappings such as:

```
u0:k10000:r10000
```

When a process creates or wants to change ownership of a file, or when the ownership of a file is read from disk by a filesystem, the userspace id is immediately translated into a kernel id according to the idmapping associated with the relevant user namespace.

For instance, consider a file that is stored on disk by a filesystem as being owned by `u1000`:

- If a filesystem were to be mounted in the initial user namespaces (as most filesystems are) then the initial idmapping will be used. As we saw this is simply the identity idmapping. This would mean id `u1000` read from disk would be mapped to id `k1000`. So an inode's `i_uid` and `i_gid` field would contain `k1000`.
- If a filesystem were to be mounted with an idmapping of `u0:k10000:r10000` then `u1000` read from disk would be mapped to `k11000`. So an inode's `i_uid` and `i_gid` would contain `k11000`.

### 1.16.3 Translation algorithms

We've already seen briefly that it is possible to translate between different idmappings. We'll now take a closer look how that works.

#### Crossmapping

This translation algorithm is used by the kernel in quite a few places. For example, it is used when reporting back the ownership of a file to userspace via the `stat()` system call family.

If we've been given `k11000` from one idmapping we can map that id up in another idmapping. In order for this to work both idmappings need to contain the same kernel id in their kernel idmapsets. For example, consider the following idmappings:

```
1. u0:k10000:r10000
2. u20000:k10000:r10000
```

and we are mapping `u1000` down to `k11000` in the first idmapping . We can then translate `k11000` into a userspace id in the second idmapping using the kernel idmapset of the second idmapping:

```
/* Map the kernel id up into a userspace id in the second idmapping. */
from_kuid(u20000:k10000:r10000, k11000) = u21000
```

Note, how we can get back to the kernel id in the first idmapping by inverting the algorithm:

```
/* Map the userspace id down into a kernel id in the second idmapping. */
make_kuid(u20000:k10000:r10000, u21000) = k11000

/* Map the kernel id up into a userspace id in the first idmapping. */
from_kuid(u0:k10000:r10000, k11000) = u1000
```

This algorithm allows us to answer the question what userspace id a given kernel id corresponds to in a given idmapping. In order to be able to answer this question both idmappings need to contain the same kernel id in their respective kernel idmapsets.

For example, when the kernel reads a raw userspace id from disk it maps it down into a kernel id according to the idmapping associated with the filesystem. Let's assume the filesystem was mounted with an idmapping of `u0:k20000:r10000` and it reads a file owned by `u1000` from disk. This means `u1000` will be mapped to `k21000` which is what will be stored in the inode's `i_uid` and `i_gid` field.

When someone in userspace calls `stat()` or a related function to get ownership information about the file the kernel can't simply map the id back up according to the filesystem's idmapping as this would give the wrong owner if the caller is using an idmapping.

So the kernel will map the id back up in the idmapping of the caller. Let's assume the caller has the somewhat unconventional idmapping `u3000:k20000:r10000` then `k21000` would map back up to `u4000`. Consequently the user would see that this file is owned by `u4000`.

### Remapping

It is possible to translate a kernel id from one idmapping to another one via the userspace idmapset of the two idmappings. This is equivalent to remapping a kernel id.

Let's look at an example. We are given the following two idmappings:

```
1. u0:k10000:r10000
2. u0:k20000:r10000
```

and we are given `k11000` in the first idmapping. In order to translate this kernel id in the first idmapping into a kernel id in the second idmapping we need to perform two steps:

1. Map the kernel id up into a userspace id in the first idmapping:

   ```
   /* Map the kernel id up into a userspace id in the first idmapping. */
   from_kuid(u0:k10000:r10000, k11000) = u1000
   ```

2. Map the userspace id down into a kernel id in the second idmapping:

   ```
   /* Map the userspace id down into a kernel id in the second idmapping. */
   make_kuid(u0:k20000:r10000, u1000) = k21000
   ```

As you can see we used the userspace idmapset in both idmappings to translate the kernel id in one idmapping to a kernel id in another idmapping.

This allows us to answer the question what kernel id we would need to use to get the same userspace id in another idmapping. In order to be able to answer this question both idmappings need to contain the same userspace id in their respective userspace idmapsets.

Note, how we can easily get back to the kernel id in the first idmapping by inverting the algorithm:

1. Map the kernel id up into a userspace id in the second idmapping:

```
/* Map the kernel id up into a userspace id in the second idmapping. */
from_kuid(u0:k20000:r10000, k21000) = u1000
```

2. Map the userspace id down into a kernel id in the first idmapping:

```
/* Map the userspace id down into a kernel id in the first idmapping. */
make_kuid(u0:k10000:r10000, u1000) = k11000
```

Another way to look at this translation is to treat it as inverting one idmapping and applying another idmapping if both idmappings have the relevant userspace id mapped. This will come in handy when working with idmapped mounts.

### Invalid translations

It is never valid to use an id in the kernel idmapset of one idmapping as the id in the userspace idmapset of another or the same idmapping. While the kernel idmapset always indicates an idmapset in the kernel id space the userspace idmapset indicates a userspace id. So the following translations are forbidden:

```
/* Map the userspace id down into a kernel id in the first idmapping. */
make_kuid(u0:k10000:r10000, u1000) = k11000

/* INVALID: Map the kernel id down into a kernel id in the second idmapping. */
make_kuid(u10000:k20000:r10000, k110000) = k21000
                                  ~~~~~~~
```

and equally wrong:

```
/* Map the kernel id up into a userspace id in the first idmapping. */
from_kuid(u0:k10000:r10000, k11000) = u1000

/* INVALID: Map the userspace id up into a userspace id in the second␣
 ↪idmapping. */
from_kuid(u20000:k0:r10000, u1000) = k21000
                             ~~~~~
```

Since userspace ids have type `uid_t` and `gid_t` and kernel ids have type `kuid_t` and `kgid_t` the compiler will throw an error when they are conflated. So the two examples above would cause a compilation failure.

### 1.16.4 Idmappings when creating filesystem objects

The concepts of mapping an id down or mapping an id up are expressed in the two kernel functions filesystem developers are rather familiar with and which we've already used in this document:

```
/* Map the userspace id down into a kernel id. */
make_kuid(idmapping, uid)

/* Map the kernel id up into a userspace id. */
from_kuid(idmapping, kuid)
```

We will take an abbreviated look into how idmappings figure into creating filesystem objects. For simplicity we will only look at what happens when the VFS has already completed path lookup right before it calls into the filesystem itself. So we're concerned with what happens when e.g. `vfs_mkdir()` is called. We will also assume that the directory we're creating filesystem objects in is readable and writable for everyone.

When creating a filesystem object the caller will look at the caller's filesystem ids. These are just regular `uid_t` and `gid_t` userspace ids but they are exclusively used when determining file ownership which is why they are called ”filesystem ids”. They are usually identical to the uid and gid of the caller but can differ. We will just assume they are always identical to not get lost in too many details.

When the caller enters the kernel two things happen:

1. Map the caller's userspace ids down into kernel ids in the caller's idmapping. (To be precise, the kernel will simply look at the kernel ids stashed in the credentials of the current task but for our education we'll pretend this translation happens just in time.)
2. Verify that the caller's kernel ids can be mapped up to userspace ids in the filesystem's idmapping.

The second step is important as regular filesystem will ultimately need to map the kernel id back up into a userspace id when writing to disk. So with the second step the kernel guarantees that a valid userspace id can be written to disk. If it can't the kernel will refuse the creation request to not even remotely risk filesystem corruption.

The astute reader will have realized that this is simply a variation of the crossmapping algorithm we mentioned above in a previous section. First, the kernel maps the caller's userspace id down into a kernel id according to the caller's idmapping and then maps that kernel id up according to the filesystem's idmapping.

From the implementation point it's worth mentioning how idmappings are represented. All idmappings are taken from the corresponding user namespace.

- caller's idmapping (usually taken from `current_user_ns()`)
- filesystem's idmapping (`sb->s_user_ns`)
- mount's idmapping (`mnt_idmap(vfsmnt)`)

Let's see some examples with caller/filesystem idmapping but without mount idmappings. This will exhibit some problems we can hit. After that we will revisit/reconsider these examples, this time using mount idmappings, to see how they can solve the problems we observed before.

### Example 1

```
caller id:            u1000
caller idmapping:     u0:k0:r4294967295
filesystem idmapping: u0:k0:r4294967295
```

Both the caller and the filesystem use the identity idmapping:

1. Map the caller's userspace ids into kernel ids in the caller's idmapping:

   ```
   make_kuid(u0:k0:r4294967295, u1000) = k1000
   ```

2. Verify that the caller's kernel ids can be mapped to userspace ids in the filesystem's idmapping.

   For this second step the kernel will call the function `fsuidgid_has_mapping()` which ultimately boils down to calling `from_kuid()`:

   ```
   from_kuid(u0:k0:r4294967295, k1000) = u1000
   ```

In this example both idmappings are the same so there's nothing exciting going on. Ultimately the userspace id that lands on disk will be `u1000`.

### Example 2

```
caller id:            u1000
caller idmapping:     u0:k10000:r10000
filesystem idmapping: u0:k20000:r10000
```

1. Map the caller's userspace ids down into kernel ids in the caller's idmapping:

   ```
   make_kuid(u0:k10000:r10000, u1000) = k11000
   ```

2. Verify that the caller's kernel ids can be mapped up to userspace ids in the filesystem's idmapping:

   ```
   from_kuid(u0:k20000:r10000, k11000) = u-1
   ```

It's immediately clear that while the caller's userspace id could be successfully mapped down into kernel ids in the caller's idmapping the kernel ids could not be mapped up according to the filesystem's idmapping. So the kernel will deny this creation request.

Note that while this example is less common, because most filesystem can't be mounted with non-initial idmappings this is a general problem as we can see in the next examples.

### Example 3

```
caller id:            u1000
caller idmapping:     u0:k10000:r10000
filesystem idmapping: u0:k0:r4294967295
```

1. Map the caller's userspace ids down into kernel ids in the caller's idmapping:

   ```
   make_kuid(u0:k10000:r10000, u1000) = k11000
   ```

2. Verify that the caller's kernel ids can be mapped up to userspace ids in the filesystem's idmapping:

   ```
   from_kuid(u0:k0:r4294967295, k11000) = u11000
   ```

We can see that the translation always succeeds. The userspace id that the filesystem will ultimately put to disk will always be identical to the value of the kernel id that was created in the caller's idmapping. This has mainly two consequences.

First, that we can't allow a caller to ultimately write to disk with another userspace id. We could only do this if we were to mount the whole filesystem with the caller's or another idmapping. But that solution is limited to a few filesystems and not very flexible. But this is a use-case that is pretty important in containerized workloads.

Second, the caller will usually not be able to create any files or access directories that have stricter permissions because none of the filesystem's kernel ids map up into valid userspace ids in the caller's idmapping

1. Map raw userspace ids down to kernel ids in the filesystem's idmapping:

   ```
   make_kuid(u0:k0:r4294967295, u1000) = k1000
   ```

2. Map kernel ids up to userspace ids in the caller's idmapping:

   ```
   from_kuid(u0:k10000:r10000, k1000) = u-1
   ```

### Example 4

```
file id:              u1000
caller idmapping:     u0:k10000:r10000
filesystem idmapping: u0:k0:r4294967295
```

In order to report ownership to userspace the kernel uses the crossmapping algorithm introduced in a previous section:

1. Map the userspace id on disk down into a kernel id in the filesystem's idmapping:

   ```
   make_kuid(u0:k0:r4294967295, u1000) = k1000
   ```

2. Map the kernel id up into a userspace id in the caller's idmapping:

   ```
   from_kuid(u0:k10000:r10000, k1000) = u-1
   ```

The crossmapping algorithm fails in this case because the kernel id in the filesystem idmapping cannot be mapped up to a userspace id in the caller's idmapping. Thus, the kernel will report the ownership of this file as the overflowid.

#### Example 5

```
file id:              u1000
caller idmapping:     u0:k10000:r10000
filesystem idmapping: u0:k20000:r10000
```

In order to report ownership to userspace the kernel uses the crossmapping algorithm introduced in a previous section:

1. Map the userspace id on disk down into a kernel id in the filesystem's idmapping:

   ```
   make_kuid(u0:k20000:r10000, u1000) = k21000
   ```

2. Map the kernel id up into a userspace id in the caller's idmapping:

   ```
   from_kuid(u0:k10000:r10000, k21000) = u-1
   ```

Again, the crossmapping algorithm fails in this case because the kernel id in the filesystem idmapping cannot be mapped to a userspace id in the caller's idmapping. Thus, the kernel will report the ownership of this file as the overflowid.

Note how in the last two examples things would be simple if the caller would be using the initial idmapping. For a filesystem mounted with the initial idmapping it would be trivial. So we only consider a filesystem with an idmapping of `u0:k20000:r10000`:

1. Map the userspace id on disk down into a kernel id in the filesystem's idmapping:

   ```
   make_kuid(u0:k20000:r10000, u1000) = k21000
   ```

2. Map the kernel id up into a userspace id in the caller's idmapping:

   ```
   from_kuid(u0:k0:r4294967295, k21000) = u21000
   ```

### 1.16.5 Idmappings on idmapped mounts

The examples we've seen in the previous section where the caller's idmapping and the filesystem's idmapping are incompatible causes various issues for workloads. For a more complex but common example, consider two containers started on the host. To completely prevent the two containers from affecting each other, an administrator may often use different non-overlapping idmappings for the two containers:

```
container1 idmapping:  u0:k10000:r10000
container2 idmapping:  u0:k20000:r10000
filesystem idmapping:  u0:k30000:r10000
```

An administrator wanting to provide easy read-write access to the following set of files:

```
dir id:       u0
dir/file1 id: u1000
dir/file2 id: u2000
```

to both containers currently can't.

Of course the administrator has the option to recursively change ownership via `chown()`. For example, they could change ownership so that `dir` and all files below it can be crossmapped from the filesystem's into the container's idmapping. Let's assume they change ownership so it is compatible with the first container's idmapping:

```
dir id:       u10000
dir/file1 id: u11000
dir/file2 id: u12000
```

This would still leave `dir` rather useless to the second container. In fact, `dir` and all files below it would continue to appear owned by the overflowid for the second container.

Or consider another increasingly popular example. Some service managers such as systemd implement a concept called ”portable home directories”. A user may want to use their home directories on different machines where they are assigned different login userspace ids. Most users will have `u1000` as the login id on their machine at home and all files in their home directory will usually be owned by `u1000`. At uni or at work they may have another login id such as `u1125`. This makes it rather difficult to interact with their home directory on their work machine.

In both cases changing ownership recursively has grave implications. The most obvious one is that ownership is changed globally and permanently. In the home directory case this change in ownership would even need to happen every time the user switches from their home to their work machine. For really large sets of files this becomes increasingly costly.

If the user is lucky, they are dealing with a filesystem that is mountable inside user namespaces. But this would also change ownership globally and the change in ownership is tied to the lifetime of the filesystem mount, i.e. the superblock. The only way to change ownership is to completely unmount the filesystem and mount it again in another user namespace. This is usually impossible because it would mean that all users currently accessing the filesystem can't anymore. And it means that `dir` still can't be shared between two containers with different idmappings. But usually the user doesn't even have this option since most filesystems aren't mountable inside containers. And not having them mountable might be desirable as it doesn't require the filesystem to deal with malicious filesystem images.

But the usecases mentioned above and more can be handled by idmapped mounts. They allow to expose the same set of dentries with different ownership at different mounts. This is achieved by marking the mounts with a user namespace through the `mount_setattr()` system call. The idmapping associated with it is then used to translate from the caller's idmapping to the filesystem's idmapping and vica versa using the remapping algorithm we introduced above.

Idmapped mounts make it possible to change ownership in a temporary and localized way. The ownership changes are restricted to a specific mount and the ownership changes are tied to the lifetime of the mount. All other users and locations where the filesystem is exposed are unaffected.

Filesystems that support idmapped mounts don't have any real reason to support being mountable inside user namespaces. A filesystem could be exposed completely under an idmapped

mount to get the same effect. This has the advantage that filesystems can leave the creation of the superblock to privileged users in the initial user namespace.

However, it is perfectly possible to combine idmapped mounts with filesystems mountable inside user namespaces. We will touch on this further below.

### Filesystem types vs idmapped mount types

With the introduction of idmapped mounts we need to distinguish between filesystem ownership and mount ownership of a VFS object such as an inode. The owner of a inode might be different when looked at from a filesystem perspective than when looked at from an idmapped mount. Such fundamental conceptual distinctions should almost always be clearly expressed in the code. So, to distinguish idmapped mount ownership from filesystem ownership separate types have been introduced.

If a uid or gid has been generated using the filesystem or caller's idmapping then we will use the `kuid_t` and `kgid_t` types. However, if a uid or gid has been generated using a mount idmapping then we will be using the dedicated `vfsuid_t` and `vfsgid_t` types.

All VFS helpers that generate or take uids and gids as arguments use the `vfsuid_t` and `vfsgid_t` types and we will be able to rely on the compiler to catch errors that originate from conflating filesystem and VFS uids and gids.

The `vfsuid_t` and `vfsgid_t` types are often mapped from and to `kuid_t` and `kgid_t` types similar how `kuid_t` and `kgid_t` types are mapped from and to `uid_t` and `gid_t` types:

```
uid_t <--> kuid_t <--> vfsuid_t
gid_t <--> kgid_t <--> vfsgid_t
```

Whenever we report ownership based on a `vfsuid_t` or `vfsgid_t` type, e.g., during `stat()`, or store ownership information in a shared VFS object based on a `vfsuid_t` or `vfsgid_t` type, e.g., during `chown()` we can use the `vfsuid_into_kuid()` and `vfsgid_into_kgid()` helpers.

To illustrate why this helper currently exists, consider what happens when we change ownership of an inode from an idmapped mount. After we generated a `vfsuid_t` or `vfsgid_t` based on the mount idmapping we later commit to this `vfsuid_t` or `vfsgid_t` to become the new filesystem wide ownership. Thus, we are turning the `vfsuid_t` or `vfsgid_t` into a global `kuid_t` or `kgid_t`. And this can be done by using `vfsuid_into_kuid()` and `vfsgid_into_kgid()`.

Note, whenever a shared VFS object, e.g., a cached `struct inode` or a cached `struct posix_acl`, stores ownership information a filesystem or "global" `kuid_t` and `kgid_t` must be used. Ownership expressed via `vfsuid_t` and `vfsgid_t` is specific to an idmapped mount.

We already noted that `vfsuid_t` and `vfsgid_t` types are generated based on mount idmappings whereas `kuid_t` and `kgid_t` types are generated based on filesystem idmappings. To prevent abusing filesystem idmappings to generate `vfsuid_t` or `vfsgid_t` types or mount idmappings to generate `kuid_t` or `kgid_t` types filesystem idmappings and mount idmappings are different types as well.

All helpers that map to or from `vfsuid_t` and `vfsgid_t` types require a mount idmapping to be passed which is of type `struct mnt_idmap`. Passing a filesystem or caller idmapping will cause a compilation error.

Similar to how we prefix all userspace ids in this document with `u` and all kernel ids with `k` we will prefix all VFS ids with `v`. So a mount idmapping will be written as: `u0:v10000:r10000`.

### Remapping helpers

Idmapping functions were added that translate between idmappings. They make use of the remapping algorithm we've introduced earlier. We're going to look at:

- `i_uid_into_vfsuid()` and `i_gid_into_vfsgid()`

  The `i_*id_into_vfs*id()` functions translate filesystem's kernel ids into VFS ids in the mount's idmapping:

```
/* Map the filesystem's kernel id up into a userspace id in the filesystem
 ↪'s idmapping. */
from_kuid(filesystem, kid) = uid

/* Map the filesystem's userspace id down ito a VFS id in the mount's␣
 ↪idmapping. */
make_kuid(mount, uid) = kuid
```

- `mapped_fsuid()` and `mapped_fsgid()`

  The `mapped_fs*id()` functions translate the caller's kernel ids into kernel ids in the filesystem's idmapping. This translation is achieved by remapping the caller's VFS ids using the mount's idmapping:

```
/* Map the caller's VFS id up into a userspace id in the mount's idmapping.
 ↪ */
from_kuid(mount, kid) = uid

/* Map the mount's userspace id down into a kernel id in the filesystem's␣
 ↪idmapping. */
make_kuid(filesystem, uid) = kuid
```

- `vfsuid_into_kuid()` and `vfsgid_into_kgid()`

  Whenever

Note that these two functions invert each other. Consider the following idmappings:

```
caller idmapping:     u0:k10000:r10000
filesystem idmapping: u0:k20000:r10000
mount idmapping:      u0:v10000:r10000
```

Assume a file owned by `u1000` is read from disk. The filesystem maps this id to `k21000` according to its idmapping. This is what is stored in the inode's `i_uid` and `i_gid` fields.

When the caller queries the ownership of this file via `stat()` the kernel would usually simply use the crossmapping algorithm and map the filesystem's kernel id up to a userspace id in the caller's idmapping.

But when the caller is accessing the file on an idmapped mount the kernel will first call `i_uid_into_vfsuid()` thereby translating the filesystem's kernel id into a VFS id in the mount's idmapping:

```
i_uid_into_vfsuid(k21000):
  /* Map the filesystem's kernel id up into a userspace id. */
```

```
  from_kuid(u0:k20000:r10000, k21000) = u1000

  /* Map the filesystem's userspace id down into a VFS id in the mount's␣
↪idmapping. */
  make_kuid(u0:v10000:r10000, u1000) = v11000
```

Finally, when the kernel reports the owner to the caller it will turn the VFS id in the mount's idmapping into a userspace id in the caller's idmapping:

```
k11000 = vfsuid_into_kuid(v11000)
from_kuid(u0:k10000:r10000, k11000) = u1000
```

We can test whether this algorithm really works by verifying what happens when we create a new file. Let's say the user is creating a file with `u1000`.

The kernel maps this to `k11000` in the caller's idmapping. Usually the kernel would now apply the crossmapping, verifying that `k11000` can be mapped to a userspace id in the filesystem's idmapping. Since `k11000` can't be mapped up in the filesystem's idmapping directly this creation request fails.

But when the caller is accessing the file on an idmapped mount the kernel will first call `mapped_fs*id()` thereby translating the caller's kernel id into a VFS id according to the mount's idmapping:

```
mapped_fsuid(k11000):
   /* Map the caller's kernel id up into a userspace id in the mount's␣
↪idmapping. */
   from_kuid(u0:k10000:r10000, k11000) = u1000

   /* Map the mount's userspace id down into a kernel id in the filesystem's␣
↪idmapping. */
   make_kuid(u0:v20000:r10000, u1000) = v21000
```

When finally writing to disk the kernel will then map `v21000` up into a userspace id in the filesystem's idmapping:

```
k21000 = vfsuid_into_kuid(v21000)
from_kuid(u0:k20000:r10000, k21000) = u1000
```

As we can see, we end up with an invertible and therefore information preserving algorithm. A file created from `u1000` on an idmapped mount will also be reported as being owned by `u1000` and vica versa.

Let's now briefly reconsider the failing examples from earlier in the context of idmapped mounts.

### Example 2 reconsidered

```
caller id:             u1000
caller idmapping:      u0:k10000:r10000
filesystem idmapping: u0:k20000:r10000
mount idmapping:       u0:v10000:r10000
```

When the caller is using a non-initial idmapping the common case is to attach the same idmapping to the mount. We now perform three steps:

1. Map the caller's userspace ids into kernel ids in the caller's idmapping:

   ```
   make_kuid(u0:k10000:r10000, u1000) = k11000
   ```

2. Translate the caller's VFS id into a kernel id in the filesystem's idmapping:

   ```
   mapped_fsuid(v11000):
     /* Map the VFS id up into a userspace id in the mount's idmapping. */
     from_kuid(u0:v10000:r10000, v11000) = u1000

     /* Map the userspace id down into a kernel id in the filesystem's␣
    ↪idmapping. */
     make_kuid(u0:k20000:r10000, u1000) = k21000
   ```

2. Verify that the caller's kernel ids can be mapped to userspace ids in the filesystem's idmapping:

   ```
   from_kuid(u0:k20000:r10000, k21000) = u1000
   ```

So the ownership that lands on disk will be `u1000`.

### Example 3 reconsidered

```
caller id:             u1000
caller idmapping:      u0:k10000:r10000
filesystem idmapping: u0:k0:r4294967295
mount idmapping:       u0:v10000:r10000
```

The same translation algorithm works with the third example.

1. Map the caller's userspace ids into kernel ids in the caller's idmapping:

   ```
   make_kuid(u0:k10000:r10000, u1000) = k11000
   ```

2. Translate the caller's VFS id into a kernel id in the filesystem's idmapping:

   ```
   mapped_fsuid(v11000):
      /* Map the VFS id up into a userspace id in the mount's idmapping. */
      from_kuid(u0:v10000:r10000, v11000) = u1000
   ```

```
  /* Map the userspace id down into a kernel id in the filesystem's␣
↪idmapping. */
  make_kuid(u0:k0:r4294967295, u1000) = k1000
```

2. Verify that the caller's kernel ids can be mapped to userspace ids in the filesystem's idmapping:

```
from_kuid(u0:k0:r4294967295, k21000) = u1000
```

So the ownership that lands on disk will be `u1000`.

### Example 4 reconsidered

```
file id:              u1000
caller idmapping:     u0:k10000:r10000
filesystem idmapping: u0:k0:r4294967295
mount idmapping:      u0:v10000:r10000
```

In order to report ownership to userspace the kernel now does three steps using the translation algorithm we introduced earlier:

1. Map the userspace id on disk down into a kernel id in the filesystem's idmapping:

```
make_kuid(u0:k0:r4294967295, u1000) = k1000
```

2. Translate the kernel id into a VFS id in the mount's idmapping:

```
i_uid_into_vfsuid(k1000):
  /* Map the kernel id up into a userspace id in the filesystem's␣
↪idmapping. */
  from_kuid(u0:k0:r4294967295, k1000) = u1000

  /* Map the userspace id down into a VFS id in the mounts's idmapping. */
  make_kuid(u0:v10000:r10000, u1000) = v11000
```

3. Map the VFS id up into a userspace id in the caller's idmapping:

```
k11000 = vfsuid_into_kuid(v11000)
from_kuid(u0:k10000:r10000, k11000) = u1000
```

Earlier, the caller's kernel id couldn't be crossmapped in the filesystems's idmapping. With the idmapped mount in place it now can be crossmapped into the filesystem's idmapping via the mount's idmapping. The file will now be created with `u1000` according to the mount's idmapping.

### Example 5 reconsidered

```
file id:              u1000
caller idmapping:     u0:k10000:r10000
filesystem idmapping: u0:k20000:r10000
mount idmapping:      u0:v10000:r10000
```

Again, in order to report ownership to userspace the kernel now does three steps using the translation algorithm we introduced earlier:

1. Map the userspace id on disk down into a kernel id in the filesystem's idmapping:

   ```
   make_kuid(u0:k20000:r10000, u1000) = k21000
   ```

2. Translate the kernel id into a VFS id in the mount's idmapping:

   ```
   i_uid_into_vfsuid(k21000):
     /* Map the kernel id up into a userspace id in the filesystem's␣
    ↪idmapping. */
     from_kuid(u0:k20000:r10000, k21000) = u1000

     /* Map the userspace id down into a VFS id in the mounts's idmapping. */
     make_kuid(u0:v10000:r10000, u1000) = v11000
   ```

3. Map the VFS id up into a userspace id in the caller's idmapping:

   ```
   k11000 = vfsuid_into_kuid(v11000)
   from_kuid(u0:k10000:r10000, k11000) = u1000
   ```

Earlier, the file's kernel id couldn't be crossmapped in the filesystems's idmapping. With the idmapped mount in place it now can be crossmapped into the filesystem's idmapping via the mount's idmapping. The file is now owned by `u1000` according to the mount's idmapping.

### Changing ownership on a home directory

We've seen above how idmapped mounts can be used to translate between idmappings when either the caller, the filesystem or both uses a non-initial idmapping. A wide range of usecases exist when the caller is using a non-initial idmapping. This mostly happens in the context of containerized workloads. The consequence is as we have seen that for both, filesystem's mounted with the initial idmapping and filesystems mounted with non-initial idmappings, access to the filesystem isn't working because the kernel ids can't be crossmapped between the caller's and the filesystem's idmapping.

As we've seen above idmapped mounts provide a solution to this by remapping the caller's or filesystem's idmapping according to the mount's idmapping.

Aside from containerized workloads, idmapped mounts have the advantage that they also work when both the caller and the filesystem use the initial idmapping which means users on the host can change the ownership of directories and files on a per-mount basis.

Consider our previous example where a user has their home directory on portable storage. At home they have id `u1000` and all files in their home directory are owned by `u1000` whereas at uni or work they have login id `u1125`.

Taking their home directory with them becomes problematic. They can't easily access their files, they might not be able to write to disk without applying lax permissions or ACLs and even if they can, they will end up with an annoying mix of files and directories owned by `u1000` and `u1125`.

Idmapped mounts allow to solve this problem. A user can create an idmapped mount for their home directory on their work computer or their computer at home depending on what ownership they would prefer to end up on the portable storage itself.

Let's assume they want all files on disk to belong to `u1000`. When the user plugs in their portable storage at their work station they can setup a job that creates an idmapped mount with the minimal idmapping `u1000:k1125:r1`. So now when they create a file the kernel performs the following steps we already know from above::

```
caller id:             u1125
caller idmapping:      u0:k0:r4294967295
filesystem idmapping: u0:k0:r4294967295
mount idmapping:       u1000:v1125:r1
```

1. Map the caller's userspace ids into kernel ids in the caller's idmapping:

   ```
   make_kuid(u0:k0:r4294967295, u1125) = k1125
   ```

2. Translate the caller's VFS id into a kernel id in the filesystem's idmapping:

   ```
   mapped_fsuid(v1125):
     /* Map the VFS id up into a userspace id in the mount's idmapping. */
     from_kuid(u1000:v1125:r1, v1125) = u1000

     /* Map the userspace id down into a kernel id in the filesystem's␣
    ↪idmapping. */
     make_kuid(u0:k0:r4294967295, u1000) = k1000
   ```

2. Verify that the caller's filesystem ids can be mapped to userspace ids in the filesystem's idmapping:

   ```
   from_kuid(u0:k0:r4294967295, k1000) = u1000
   ```

So ultimately the file will be created with `u1000` on disk.

Now let's briefly look at what ownership the caller with id `u1125` will see on their work computer:

```
file id:               u1000
caller idmapping:      u0:k0:r4294967295
filesystem idmapping: u0:k0:r4294967295
mount idmapping:       u1000:v1125:r1
```

1. Map the userspace id on disk down into a kernel id in the filesystem's idmapping:

   ```
   make_kuid(u0:k0:r4294967295, u1000) = k1000
   ```

2. Translate the kernel id into a VFS id in the mount's idmapping:

```
i_uid_into_vfsuid(k1000):
  /* Map the kernel id up into a userspace id in the filesystem's␣
↪idmapping. */
  from_kuid(u0:k0:r4294967295, k1000) = u1000

  /* Map the userspace id down into a VFS id in the mounts's idmapping. */
  make_kuid(u1000:v1125:r1, u1000) = v1125
```

3. Map the VFS id up into a userspace id in the caller's idmapping:

```
k1125 = vfsuid_into_kuid(v1125)
from_kuid(u0:k0:r4294967295, k1125) = u1125
```

So ultimately the caller will be reported that the file belongs to `u1125` which is the caller's userspace id on their workstation in our example.

The raw userspace id that is put on disk is `u1000` so when the user takes their home directory back to their home computer where they are assigned `u1000` using the initial idmapping and mount the filesystem with the initial idmapping they will see all those files owned by `u1000`.

## 1.17 Automount Support

Support is available for filesystems that wish to do automounting support (such as kAFS which can be found in fs/afs/ and NFS in fs/nfs/). This facility includes allowing in-kernel mounts to be performed and mountpoint degradation to be requested. The latter can also be requested by userspace.

### 1.17.1 In-Kernel Automounting

See section "Mount Traps" of *autofs - how it works*

Then from userspace, you can just do something like:

```
[root@andromeda root]# mount -t afs \#root.afs. /afs
[root@andromeda root]# ls /afs
asd  cambridge  cambridge.redhat.com  grand.central.org
[root@andromeda root]# ls /afs/cambridge
afsdoc
[root@andromeda root]# ls /afs/cambridge/afsdoc/
ChangeLog  html  LICENSE  pdf  RELNOTES-1.2.2
```

And then if you look in the mountpoint catalogue, you'll see something like:

```
[root@andromeda root]# cat /proc/mounts
...
#root.afs. /afs afs rw 0 0
#root.cell. /afs/cambridge.redhat.com afs rw 0 0
#afsdoc. /afs/cambridge.redhat.com/afsdoc afs rw 0 0
```

### 1.17.2 Automatic Mountpoint Expiry

Automatic expiration of mountpoints is easy, provided you've mounted the mountpoint to be expired in the automounting procedure outlined separately.

To do expiration, you need to follow these steps:

(1) Create at least one list off which the vfsmounts to be expired can be hung.

(2) When a new mountpoint is created in the ->d_automount method, add the mnt to the list using *mnt_set_expiry()*:

```
mnt_set_expiry(newmnt, &afs_vfsmounts);
```

(3) When you want mountpoints to be expired, call mark_mounts_for_expiry() with a pointer to this list. This will process the list, marking every vfsmount thereon for potential expiry on the next call.

If a vfsmount was already flagged for expiry, and if its usage count is 1 (it's only referenced by its parent vfsmount), then it will be deleted from the namespace and thrown away (effectively unmounted).

It may prove simplest to simply call this at regular intervals, using some sort of timed event to drive it.

The expiration flag is cleared by calls to mntput. This means that expiration will only happen on the second expiration request after the last time the mountpoint was accessed.

If a mountpoint is moved, it gets removed from the expiration list. If a bind mount is made on an expirable mount, the new vfsmount will not be on the expiration list and will not expire.

If a namespace is copied, all mountpoints contained therein will be copied, and the copies of those that are on an expiration list will be added to the same expiration list.

### 1.17.3 Userspace Driven Expiry

As an alternative, it is possible for userspace to request expiry of any mountpoint (though some will be rejected - the current process's idea of the rootfs for example). It does this by passing the MNT_EXPIRE flag to umount(). This flag is considered incompatible with MNT_FORCE and MNT_DETACH.

If the mountpoint in question is in referenced by something other than umount() or its parent mountpoint, an EBUSY error will be returned and the mountpoint will not be marked for expiration or unmounted.

If the mountpoint was not already marked for expiry at that time, an EAGAIN error will be given and it won't be unmounted.

Otherwise if it was already marked and it wasn't referenced, unmounting will take place as usual.

Again, the expiration flag is cleared every time anything other than umount() looks at a mountpoint.

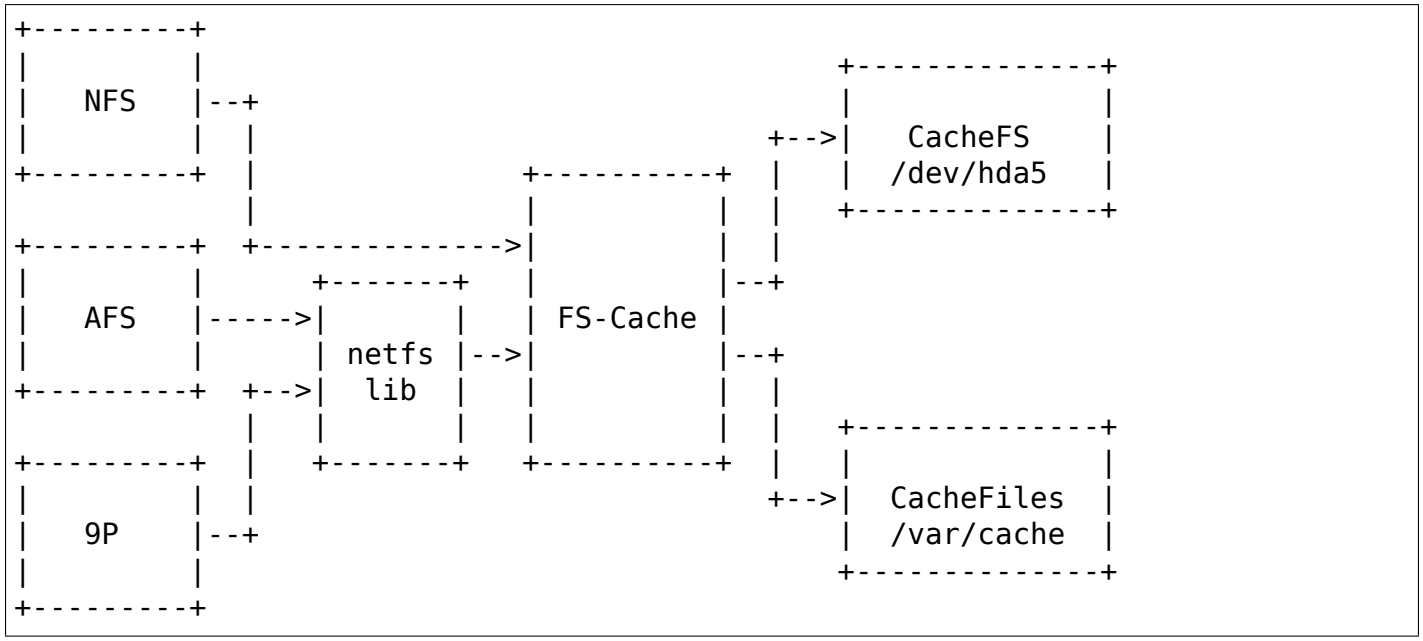
# 1.18 Filesystem Caching

## 1.18.1 General Filesystem Caching

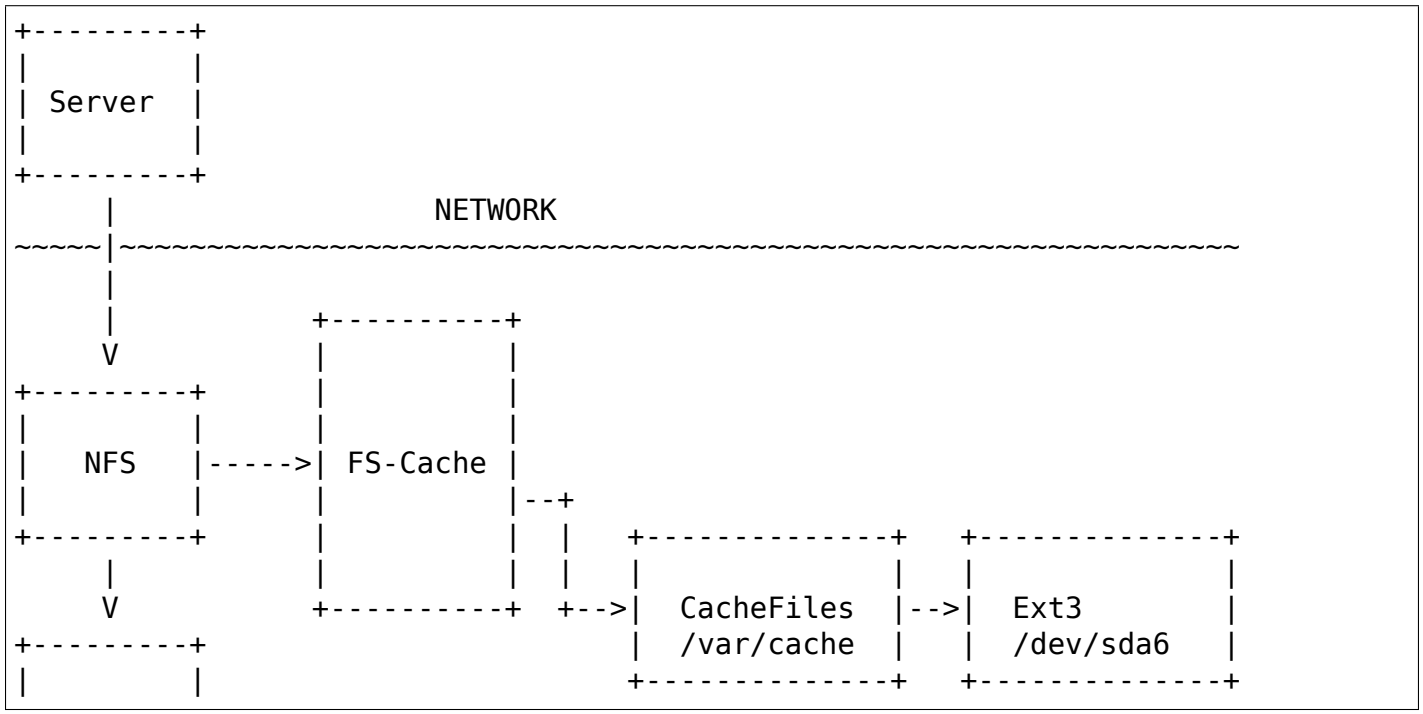
### Overview

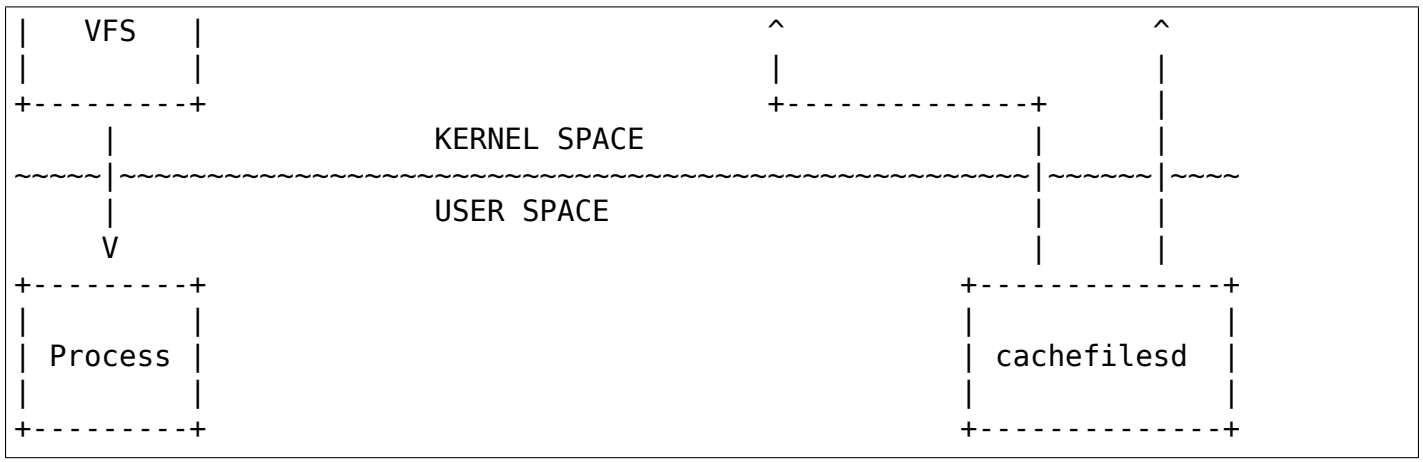
This facility is a general purpose cache for network filesystems, though it could be used for caching other things such as ISO9660 filesystems too.

FS-Cache mediates between cache backends (such as CacheFiles) and network filesystems:




Or to look at it another way, FS-Cache is a module that provides a caching facility to a network filesystem such that the cache is transparent to the user:






FS-Cache does not follow the idea of completely loading every netfs file opened in its entirety into a cache before permitting it to be accessed and then serving the pages out of that cache rather than the netfs inode because:

(1) It must be practical to operate without a cache.

(2) The size of any accessible file must not be limited to the size of the cache.

(3) The combined size of all opened files (this includes mapped libraries) must not be limited to the size of the cache.

(4) The user should not be forced to download an entire file just to do a one-off access of a small portion of it (such as might be done with the "file" program).

It instead serves the cache out in chunks as and when requested by the netfs using it.

FS-Cache provides the following facilities:

- More than one cache can be used at once. Caches can be selected explicitly by use of tags.
- Caches can be added / removed at any time, even whilst being accessed.
- The netfs is provided with an interface that allows either party to withdraw caching facilities from a file (required for (2)).
- The interface to the netfs returns as few errors as possible, preferring rather to let the netfs remain oblivious.
- There are three types of cookie: cache, volume and data file cookies. Cache cookies represent the cache as a whole and are not normally visible to the netfs; the netfs gets a volume cookie to represent a collection of files (typically something that a netfs would get for a superblock); and data file cookies are used to cache data (something that would be got for an inode).
- Volumes are matched using a key. This is a printable string that is used to encode all the information that might be needed to distinguish one superblock, say, from another. This would be a compound of things like cell name or server address, volume name or share path. It must be a valid pathname.
- Cookies are matched using a key. This is a binary blob and is used to represent the object within a volume (so the volume key need not form part of the blob). This might include things like an inode number and uniquifier or a file handle.
- Cookie resources are set up and pinned by marking the cookie in-use. This prevents the backing resources from being culled. Timed garbage collection is employed to eliminate

cookies that haven't been used for a short while, thereby reducing resource overload. This is intended to be used when a file is opened or closed.

A cookie can be marked in-use multiple times simultaneously; each mark must be unused.

- Begin/end access functions are provided to delay cache withdrawal for the duration of an operation and prevent structs from being freed whilst we're looking at them.
- Data I/O is done by asynchronous DIO to/from a buffer described by the netfs using an iov_iter.
- An invalidation facility is available to discard data from the cache and to deal with I/O that's in progress that is accessing old data.
- Cookies can be ”retired” upon release, thereby causing the object to be removed from the cache.

The netfs API to FS-Cache can be found in:

*Network Filesystem Caching API*

The cache backend API to FS-Cache can be found in:

*Cache Backend API*

### Statistical Information

If FS-Cache is compiled with the following options enabled:

```
CONFIG_FSCACHE_STATS=y
```

then it will gather certain statistics and display them through:

/proc/fs/fscache/stats

This shows counts of a number of events that can happen in FS-Cache:

| CLASS | EVENT | MEANING |
|---|---|---|
| Cookies | n=N | Number of data storage cookies allocated |
| | v=N | Number of volume index cookies allocated |
| | vcol=N | Number of volume index key collisions |
| | voom=N | Number of OOM events when allocating volume cookies |
| Acquire | n=N | Number of acquire cookie requests seen |
| | ok=N | Number of acq reqs succeeded |
| | oom=N | Number of acq reqs failed on ENOMEM |
| LRU | n=N | Number of cookies currently on the LRU |
| | exp=N | Number of cookies expired off of the LRU |
| | rmv=N | Number of cookies removed from the LRU |
| | drp=N | Number of LRU'd cookies relinquished/withdrawn |
| | at=N | Time till next LRU cull (jiffies) |
| Invals | n=N | Number of invalidations |
| Updates | n=N | Number of update cookie requests seen |
| | rsz=N | Number of resize requests |
| | rsn=N | Number of skipped resize requests |
| Relinqs | n=N | Number of relinquish cookie requests seen |
| | rtr=N | Number of rlq reqs with retire=true |
| | drop=N | Number of cookies no longer blocking re-acquisition |
| NoSpace | nwr=N | Number of write requests refused due to lack of space |
| | ncr=N | Number of create requests refused due to lack of space |
| | cull=N | Number of objects culled to make space |
| IO | rd=N | Number of read operations in the cache |
| | wr=N | Number of write operations in the cache |

Netfslib will also add some stats counters of its own.

### Cache List

FS-Cache provides a list of cache cookies:

```
/proc/fs/fscache/cookies
```

This will look something like:

```
# cat /proc/fs/fscache/caches
CACHE    REF   VOLS  OBJS  ACCES S NAME
======== ===== ===== ===== ===== = ===============
00000001     2     1  2123     1 A default
```

where the columns are:

| COLUMN | DESCRIPTION |
|---|---|
| CACHE | Cache cookie debug ID (also appears in traces) |
| REF | Number of references on the cache cookie |
| VOLS | Number of volumes cookies in this cache |
| OBJS | Number of cache objects in use |
| ACCES | Number of accesses pinning the cache |
| S | State |
| NAME | Name of the cache. |

The state can be (-) Inactive, (P)reparing, (A)ctive, (E)rror or (W)ithdrawing.

### Volume List

FS-Cache provides a list of volume cookies:

    /proc/fs/fscache/volumes

This will look something like:

```
VOLUME   REF   nCOOK ACC FL CACHE           KEY
======== ===== ===== === == =============== ================
00000001    55    54   1 00 default         afs,example.com,100058
```

where the columns are:

| COLUMN | DESCRIPTION |
|---|---|
| VOLUME | The volume cookie debug ID (also appears in traces) |
| REF | Number of references on the volume cookie |
| nCOOK | Number of cookies in the volume |
| ACC | Number of accesses pinning the cache |
| FL | Flags on the volume cookie |
| CACHE | Name of the cache or "-" |
| KEY | The indexing key for the volume |

### Cookie List

FS-Cache provides a list of cookies:

    /proc/fs/fscache/cookies

This will look something like:

```
# head /proc/fs/fscache/cookies
COOKIE   VOLUME   REF ACT ACC S FL DEF
======== ======== === === === = == ================
00000435 00000001   1   0  -1 - 08 0000000201d0800700000000000000000,␣
 ↪0000000000000000
00000436 00000001   1   0  -1 - 00 0000005601d0800800000000000000000,␣
 ↪0000000000000051
00000437 00000001   1   0  -1 - 08 00023b3001d0823f0000000000000000,␣
 ↪0000000000000000
00000438 00000001   1   0  -1 - 08 0000005801d0807b0000000000000000,␣
 ↪0000000000000000
00000439 00000001   1   0  -1 - 08 00023b3201d080a10000000000000000,␣
 ↪0000000000000000
0000043a 00000001   1   0  -1 - 08 00023b3401d080a30000000000000000,␣
 ↪0000000000000000
0000043b 00000001   1   0  -1 - 08 00023b3601d080b30000000000000000,␣
 ↪0000000000000000
0000043c 00000001   1   0  -1 - 08 00023b3801d080b40000000000000000,␣
 ↪0000000000000000
```

where the columns are:

| COLUMN | DESCRIPTION |
|---|---|
| COOKIE | The cookie debug ID (also appears in traces) |
| VOLUME | The parent volume cookie debug ID |
| REF | Number of references on the volume cookie |
| ACT | Number of times the cookie is marked for in use |
| ACC | Number of access pins in the cookie |
| S | State of the cookie |
| FL | Flags on the cookie |
| DEF | Key, auxiliary data |

### Debugging

If CONFIG_FSCACHE_DEBUG is enabled, the FS-Cache facility can have runtime debugging enabled by adjusting the value in:

```
/sys/module/fscache/parameters/debug
```

This is a bitmask of debugging streams to enable:

| BIT | VALUE | STREAM | POINT |
|---|---|---|---|
| 0 | 1 | Cache management | Function entry trace |
| 1 | 2 | | Function exit trace |
| 2 | 4 | | General |
| 3 | 8 | Cookie management | Function entry trace |
| 4 | 16 | | Function exit trace |
| 5 | 32 | | General |
| 6-8 | | | (Not used) |
| 9 | 512 | I/O operation management | Function entry trace |
| 10 | 1024 | | Function exit trace |
| 11 | 2048 | | General |

The appropriate set of values should be OR'd together and the result written to the control file. For example:

```
echo $((1|8|512)) >/sys/module/fscache/parameters/debug
```

will turn on all function entry debugging.

### 1.18.2 Network Filesystem Caching API

Fscache provides an API by which a network filesystem can make use of local caching facilities. The API is arranged around a number of principles:

(1) A cache is logically organised into volumes and data storage objects within those volumes.

(2) Volumes and data storage objects are represented by various types of cookie.

(3) Cookies have keys that distinguish them from their peers.

(4) Cookies have coherency data that allows a cache to determine if the cached data is still valid.

(5) I/O is done asynchronously where possible.

This API is used by:

```
#include <linux/fscache.h>.
```

#### Overview

The fscache hierarchy is organised on two levels from a network filesystem's point of view. The upper level represents ”volumes” and the lower level represents ”data storage objects”. These are represented by two types of cookie, hereafter referred to as ”volume cookies” and ”cookies”.

A network filesystem acquires a volume cookie for a volume using a volume key, which represents all the information that defines that volume (e.g. cell name or server address, volume ID or share name). This must be rendered as a printable string that can be used as a directory name (ie. no '/' characters and shouldn't begin with a '.'). The maximum name length is one less than the maximum size of a filename component (allowing the cache backend one char for its own purposes).

A filesystem would typically have a volume cookie for each superblock.

The filesystem then acquires a cookie for each file within that volume using an object key. Object keys are binary blobs and only need to be unique within their parent volume. The cache backend is responsible for rendering the binary blob into something it can use and may employ hash tables, trees or whatever to improve its ability to find an object. This is transparent to the network filesystem.

A filesystem would typically have a cookie for each inode, and would acquire it in iget and relinquish it when evicting the cookie.

Once it has a cookie, the filesystem needs to mark the cookie as being in use. This causes fscache to send the cache backend off to look up/create resources for the cookie in the background, to check its coherency and, if necessary, to mark the object as being under modification.

A filesystem would typically ”use” the cookie in its file open routine and unuse it in file release and it needs to use the cookie around calls to truncate the cookie locally. It *also* needs to use the cookie when the pagecache becomes dirty and unuse it when writeback is complete. This is slightly tricky, and provision is made for it.

When performing a read, write or resize on a cookie, the filesystem must first begin an operation. This copies the resources into a holding struct and puts extra pins into the cache to stop cache withdrawal from tearing down the structures being used. The actual operation can then be issued and conflicting invalidations can be detected upon completion.

The filesystem is expected to use netfslib to access the cache, but that's not actually required and it can use the fscache I/O API directly.

### Volume Registration

The first step for a network filesystem is to acquire a volume cookie for the volume it wants to access:

```
struct fscache_volume *
fscache_acquire_volume(const char *volume_key,
                       const char *cache_name,
                       const void *coherency_data,
                       size_t coherency_len);
```

This function creates a volume cookie with the specified volume key as its name and notes the coherency data.

The volume key must be a printable string with no '/' characters in it. It should begin with the name of the filesystem and should be no longer than 254 characters. It should uniquely represent the volume and will be matched with what's stored in the cache.

The caller may also specify the name of the cache to use. If specified, fscache will look up or create a cache cookie of that name and will use a cache of that name if it is online or comes online. If no cache name is specified, it will use the first cache that comes to hand and set the name to that.

The specified coherency data is stored in the cookie and will be matched against coherency data stored on disk. The data pointer may be NULL if no data is provided. If the coherency data doesn't match, the entire cache volume will be invalidated.

This function can return errors such as EBUSY if the volume key is already in use by an acquired volume or ENOMEM if an allocation failure occurred. It may also return a NULL volume cookie if fscache is not enabled. It is safe to pass a NULL cookie to any function that takes a volume cookie. This will cause that function to do nothing.

When the network filesystem has finished with a volume, it should relinquish it by calling:

```
void fscache_relinquish_volume(struct fscache_volume *volume,
                               const void *coherency_data,
                               bool invalidate);
```

This will cause the volume to be committed or removed, and if sealed the coherency data will be set to the value supplied. The amount of coherency data must match the length specified when the volume was acquired. Note that all data cookies obtained in this volume must be relinquished before the volume is relinquished.

### Data File Registration

Once it has a volume cookie, a network filesystem can use it to acquire a cookie for data storage:

```
struct fscache_cookie *
fscache_acquire_cookie(struct fscache_volume *volume,
                       u8 advice,
                       const void *index_key,
                       size_t index_key_len,
                       const void *aux_data,
                       size_t aux_data_len,
                       loff_t object_size)
```

This creates the cookie in the volume using the specified index key. The index key is a binary blob of the given length and must be unique for the volume. This is saved into the cookie. There are no restrictions on the content, but its length shouldn't exceed about three quarters of the maximum filename length to allow for encoding.

The caller should also pass in a piece of coherency data in aux_data. A buffer of size aux_data_len will be allocated and the coherency data copied in. It is assumed that the size is invariant over time. The coherency data is used to check the validity of data in the cache. Functions are provided by which the coherency data can be updated.

The file size of the object being cached should also be provided. This may be used to trim the data and will be stored with the coherency data.

This function never returns an error, though it may return a NULL cookie on allocation failure or if fscache is not enabled. It is safe to pass in a NULL volume cookie and pass the NULL cookie returned to any function that takes it. This will cause that function to do nothing.

When the network filesystem has finished with a cookie, it should relinquish it by calling:

```
void fscache_relinquish_cookie(struct fscache_cookie *cookie,
                               bool retire);
```

This will cause fscache to either commit the storage backing the cookie or delete it.

### Marking A Cookie In-Use

Once a cookie has been acquired by a network filesystem, the filesystem should tell fscache when it intends to use the cookie (typically done on file open) and should say when it has finished with it (typically on file close):

```
void fscache_use_cookie(struct fscache_cookie *cookie,
                        bool will_modify);
void fscache_unuse_cookie(struct fscache_cookie *cookie,
                          const void *aux_data,
                          const loff_t *object_size);
```

The *use* function tells fscache that it will use the cookie and, additionally, indicate if the user is intending to modify the contents locally. If not yet done, this will trigger the cache backend to go and gather the resources it needs to access/store data in the cache. This is done in the background, and so may not be complete by the time the function returns.

The *unuse* function indicates that a filesystem has finished using a cookie. It optionally updates the stored coherency data and object size and then decreases the in-use counter. When the last user unuses the cookie, it is scheduled for garbage collection. If not reused within a short time, the resources will be released to reduce system resource consumption.

A cookie must be marked in-use before it can be accessed for read, write or resize - and an in-use mark must be kept whilst there is dirty data in the pagecache in order to avoid an oops due to trying to open a file during process exit.

Note that in-use marks are cumulative. For each time a cookie is marked in-use, it must be unused.

### Resizing A Data File (Truncation)

If a network filesystem file is resized locally by truncation, the following should be called to notify the cache:

```
void fscache_resize_cookie(struct fscache_cookie *cookie,
                           loff_t new_size);
```

The caller must have first marked the cookie in-use. The cookie and the new size are passed in and the cache is synchronously resized. This is expected to be called from `->setattr()` inode operation under the inode lock.

### Data I/O API

To do data I/O operations directly through a cookie, the following functions are available:

```
int fscache_begin_read_operation(struct netfs_cache_resources *cres,
                                 struct fscache_cookie *cookie);
int fscache_read(struct netfs_cache_resources *cres,
                 loff_t start_pos,
                 struct iov_iter *iter,
                 enum netfs_read_from_hole read_hole,
                 netfs_io_terminated_t term_func,
                 void *term_func_priv);
int fscache_write(struct netfs_cache_resources *cres,
                  loff_t start_pos,
                  struct iov_iter *iter,
                  netfs_io_terminated_t term_func,
                  void *term_func_priv);
```

The *begin* function sets up an operation, attaching the resources required to the cache resources block from the cookie. Assuming it doesn't return an error (for instance, it will return -ENOBUFS if given a NULL cookie, but otherwise do nothing), then one of the other two functions can be issued.

The *read* and *write* functions initiate a direct-IO operation. Both take the previously set up cache resources block, an indication of the start file position, and an I/O iterator that describes buffer and indicates the amount of data.

The read function also takes a parameter to indicate how it should handle a partially populated region (a hole) in the disk content. This may be to ignore it, skip over an initial hole and place zeros in the buffer or give an error.

The read and write functions can be given an optional termination function that will be run on completion:

```
typedef
void (*netfs_io_terminated_t)(void *priv, ssize_t transferred_or_error,
                              bool was_async);
```

If a termination function is given, the operation will be run asynchronously and the termination function will be called upon completion. If not given, the operation will be run synchronously. Note that in the asynchronous case, it is possible for the operation to complete before the function returns.

Both the read and write functions end the operation when they complete, detaching any pinned resources.

The read operation will fail with ESTALE if invalidation occurred whilst the operation was ongoing.

### Data File Coherency

To request an update of the coherency data and file size on a cookie, the following should be called:

```
void fscache_update_cookie(struct fscache_cookie *cookie,
                           const void *aux_data,
                           const loff_t *object_size);
```

This will update the cookie's coherency data and/or file size.

### Data File Invalidation

Sometimes it will be necessary to invalidate an object that contains data. Typically this will be necessary when the server informs the network filesystem of a remote third-party change - at which point the filesystem has to throw away the state and cached data that it had for an file and reload from the server.

To indicate that a cache object should be invalidated, the following should be called:

```
void fscache_invalidate(struct fscache_cookie *cookie,
                        const void *aux_data,
                        loff_t size,
                        unsigned int flags);
```

This increases the invalidation counter in the cookie to cause outstanding reads to fail with -ESTALE, sets the coherency data and file size from the information supplied, blocks new I/O on the cookie and dispatches the cache to go and get rid of the old data.

Invalidation runs asynchronously in a worker thread so that it doesn't block too much.

### Write-Back Resource Management

To write data to the cache from network filesystem writeback, the cache resources required need to be pinned at the point the modification is made (for instance when the page is marked dirty) as it's not possible to open a file in a thread that's exiting.

The following facilities are provided to manage this:

- An inode flag, `I_PINNING_FSCACHE_WB`, is provided to indicate that an in-use is held on the cookie for this inode. It can only be changed if the the inode lock is held.
- A flag, `unpinned_fscache_wb` is placed in the `writeback_control` struct that gets set if `__writeback_single_inode()` clears `I_PINNING_FSCACHE_WB` because all the dirty pages were cleared.

To support this, the following functions are provided:

```
bool fscache_dirty_folio(struct address_space *mapping,
                         struct folio *folio,
                         struct fscache_cookie *cookie);
void fscache_unpin_writeback(struct writeback_control *wbc,
                             struct fscache_cookie *cookie);
void fscache_clear_inode_writeback(struct fscache_cookie *cookie,
                                   struct inode *inode,
                                   const void *aux);
```

The *set* function is intended to be called from the filesystem's `dirty_folio` address space operation. If `I_PINNING_FSCACHE_WB` is not set, it sets that flag and increments the use count on the cookie (the caller must already have called `fscache_use_cookie()`).

The *unpin* function is intended to be called from the filesystem's `write_inode` superblock operation. It cleans up after writing by unusing the cookie if unpinned_fscache_wb is set in the writeback_control struct.

The *clear* function is intended to be called from the netfs's `evict_inode` superblock operation. It must be called *after* `truncate_inode_pages_final()`, but *before* `clear_inode()`. This cleans up any hanging `I_PINNING_FSCACHE_WB`. It also allows the coherency data to be updated.

### Caching of Local Modifications

If a network filesystem has locally modified data that it wants to write to the cache, it needs to mark the pages to indicate that a write is in progress, and if the mark is already present, it needs to wait for it to be removed first (presumably due to an already in-progress operation). This prevents multiple competing DIO writes to the same storage in the cache.

Firstly, the netfs should determine if caching is available by doing something like:

```
bool caching = fscache_cookie_enabled(cookie);
```

If caching is to be attempted, pages should be waited for and then marked using the following functions provided by the netfs helper library:

```
void set_page_fscache(struct page *page);
void wait_on_page_fscache(struct page *page);
int wait_on_page_fscache_killable(struct page *page);
```

Once all the pages in the span are marked, the netfs can ask fscache to schedule a write of that region:

```
void fscache_write_to_cache(struct fscache_cookie *cookie,
                            struct address_space *mapping,
                            loff_t start, size_t len, loff_t i_size,
                            netfs_io_terminated_t term_func,
                            void *term_func_priv,
                            bool caching)
```

And if an error occurs before that point is reached, the marks can be removed by calling:

```
void fscache_clear_page_bits(struct address_space *mapping,
                             loff_t start, size_t len,
                             bool caching)
```

In these functions, a pointer to the mapping to which the source pages are attached is passed in and start and len indicate the size of the region that's going to be written (it doesn't have to align to page boundaries necessarily, but it does have to align to DIO boundaries on the backing filesystem). The caching parameter indicates if caching should be skipped, and if false, the functions do nothing.

The write function takes some additional parameters: the cookie representing the cache object to be written to, i_size indicates the size of the netfs file and term_func indicates an optional completion function, to which term_func_priv will be passed, along with the error or amount written.

Note that the write function will always run asynchronously and will unmark all the pages upon completion before calling term_func.

### Page Release and Invalidation

Fscache keeps track of whether we have any data in the cache yet for a cache object we've just created. It knows it doesn't have to do any reading until it has done a write and then the page it wrote from has been released by the VM, after which it *has* to look in the cache.

To inform fscache that a page might now be in the cache, the following function should be called from the `release_folio` address space op:

```
void fscache_note_page_release(struct fscache_cookie *cookie);
```

if the page has been released (ie. release_folio returned true).

Page release and page invalidation should also wait for any mark left on the page to say that a DIO write is underway from that page:

```
void wait_on_page_fscache(struct page *page);
int wait_on_page_fscache_killable(struct page *page);
```

### API Function Reference

struct fscache_volume *__fscache_acquire_volume__(const char *volume_key, const char *cache_name, const void *coherency_data, size_t coherency_len)

Register a volume as desiring caching services

**Parameters**

**const char *volume_key**
An identification string for the volume

**const char *cache_name**
The name of the cache to use (or NULL for the default)

**const void *coherency_data**
Piece of arbitrary coherency data to check (or NULL)

**size_t coherency_len**
The size of the coherency data

**Description**

Register a volume as desiring caching services if they're available. The caller must provide an identifier for the volume and may also indicate which cache it should be in. If a preexisting volume entry is found in the cache, the coherency data must match otherwise the entry will be invalidated.

Returns a cookie pointer on success, -ENOMEM if out of memory or -EBUSY if a cache volume of that name is already acquired. Note that "NULL" is a valid cookie pointer and can be returned if caching is refused.

void __fscache_relinquish_volume__(struct fscache_volume *volume, const void *coherency_data, bool invalidate)

Cease caching a volume

**Parameters**

**struct fscache_volume *volume**
The volume cookie

**const void *coherency_data**
Piece of arbitrary coherency data to set (or NULL)

**bool invalidate**
True if the volume should be invalidated

**Description**

Indicate that a filesystem no longer desires caching services for a volume. The caller must have relinquished all file cookies prior to calling this. The stored coherency data is updated.

struct fscache_cookie *__fscache_acquire_cookie__(struct fscache_volume *volume, u8 advice, const void *index_key, size_t index_key_len, const void *aux_data, size_t aux_data_len, loff_t object_size)

Acquire a cookie to represent a cache object

**Parameters**

**struct fscache_volume *volume**
The volume in which to locate/create this cookie

**u8 advice**
Advice flags (FSCACHE_COOKIE_ADV_*)

**const void *index_key**
The index key for this cookie

**size_t index_key_len**
Size of the index key

**const void *aux_data**
The auxiliary data for the cookie (may be NULL)

**size_t aux_data_len**
Size of the auxiliary data buffer

**loff_t object_size**
The initial size of object

**Description**

Acquire a cookie to represent a data file within the given cache volume.

See *Network Filesystem Caching API* for a complete description.

void **fscache_use_cookie**(struct fscache_cookie *cookie, bool will_modify)
Request usage of cookie attached to an object

**Parameters**

**struct fscache_cookie *cookie**
The cookie representing the cache object

**bool will_modify**
If cache is expected to be modified locally

**Description**

Request usage of the cookie attached to an object. The caller should tell the cache if the object's contents are about to be modified locally and then the cache can apply the policy that has been set to handle this case.

void **fscache_unuse_cookie**(struct fscache_cookie *cookie, const void *aux_data, const loff_t *object_size)
Cease usage of cookie attached to an object

**Parameters**

**struct fscache_cookie *cookie**
The cookie representing the cache object

**const void *aux_data**
Updated auxiliary data (or NULL)

**const loff_t *object_size**
Revised size of the object (or NULL)

**Description**

Cease usage of the cookie attached to an object. When the users count reaches zero then the cookie relinquishment will be permitted to proceed.

void **fscache_relinquish_cookie**(struct fscache_cookie *cookie, bool retire)

Return the cookie to the cache, maybe discarding it

**Parameters**

**struct fscache_cookie *cookie**
The cookie being returned

**bool retire**
True if the cache object the cookie represents is to be discarded

**Description**

This function returns a cookie to the cache, forcibly discarding the associated cache object if retire is set to true.

See *Network Filesystem Caching API* for a complete description.

void **fscache_update_cookie**(struct fscache_cookie *cookie, const void *aux_data, const loff_t *object_size)

Request that a cache object be updated

**Parameters**

**struct fscache_cookie *cookie**
The cookie representing the cache object

**const void *aux_data**
The updated auxiliary data for the cookie (may be NULL)

**const loff_t *object_size**
The current size of the object (may be NULL)

**Description**

Request an update of the index data for the cache object associated with the cookie. The auxiliary data on the cookie will be updated first if **aux_data** is set and the object size will be updated and the object possibly trimmed if **object_size** is set.

See *Network Filesystem Caching API* for a complete description.

void **fscache_resize_cookie**(struct fscache_cookie *cookie, loff_t new_size)

Request that a cache object be resized

**Parameters**

**struct fscache_cookie *cookie**
The cookie representing the cache object

**loff_t new_size**
The new size of the object (may be NULL)

**Description**

Request that the size of an object be changed.

See *Network Filesystem Caching API* for a complete description.

void **fscache_invalidate**(struct fscache_cookie *cookie, const void *aux_data, loff_t size, unsigned int flags)

Notify cache that an object needs invalidation

**Parameters**

**struct fscache_cookie *cookie**
The cookie representing the cache object

**const void *aux_data**
The updated auxiliary data for the cookie (may be NULL)

**loff_t size**
The revised size of the object.

**unsigned int flags**
Invalidation flags (FSCACHE_INVAL_*)

**Description**

Notify the cache that an object is needs to be invalidated and that it should abort any retrievals or stores it is doing on the cache. This increments inval_counter on the cookie which can be used by the caller to reconsider I/O requests as they complete.

If **flags** has FSCACHE_INVAL_DIO_WRITE set, this indicates that this is due to a direct I/O write and will cause caching to be disabled on this cookie until it is completely unused.

See *Network Filesystem Caching API* for a complete description.

const struct netfs_cache_ops *__fscache_operation_valid__(const struct netfs_cache_resources *cres)

Return true if operations resources are usable

**Parameters**

**const struct netfs_cache_resources *cres**
The resources to check.

**Description**

Returns a pointer to the operations table if usable or NULL if not.

int **fscache_begin_read_operation**(struct netfs_cache_resources *cres, struct fscache_cookie *cookie)

Begin a read operation for the netfs lib

**Parameters**

**struct netfs_cache_resources *cres**
The cache resources for the read being performed

**struct fscache_cookie *cookie**
The cookie representing the cache object

**Description**

Begin a read operation on behalf of the netfs helper library. **cres** indicates the cache resources to which the operation state should be attached; **cookie** indicates the cache object that will be accessed.

**cres->inval_counter** is set from **cookie->inval_counter** for comparison at the end of the operation. This allows invalidation during the operation to be detected by the caller.

**Return**

- 0 - Success
- **-ENOBUFS** – No caching available
- Other error code from the cache, such as -ENOMEM.

void **fscache_end_operation**(struct netfs_cache_resources *cres)

End the read operation for the netfs lib

**Parameters**

**struct netfs_cache_resources *cres**
The cache resources for the read operation

**Description**

Clean up the resources at the end of the read request.

int **fscache_read**(struct netfs_cache_resources *cres, loff_t start_pos, struct iov_iter *iter, enum netfs_read_from_hole read_hole, netfs_io_terminated_t term_func, void *term_func_priv)

Start a read from the cache.

**Parameters**

**struct netfs_cache_resources *cres**
The cache resources to use

**loff_t start_pos**
The beginning file offset in the cache file

**struct iov_iter *iter**
The buffer to fill - and also the length

**enum netfs_read_from_hole read_hole**
How to handle a hole in the data.

**netfs_io_terminated_t term_func**
The function to call upon completion

**void *term_func_priv**
The private data for **term_func**

**Description**

Start a read from the cache. **cres** indicates the cache object to read from and must be obtained by a call to fscache_begin_operation() beforehand.

The data is read into the iterator, **iter**, and that also indicates the size of the operation. **start_pos** is the start position in the file, though if **seek_data** is set appropriately, the cache can use SEEK_DATA to find the next piece of data, writing zeros for the hole into the iterator.

Upon termination of the operation, **term_func** will be called and supplied with **term_func_priv** plus the amount of data written, if successful, or the error code otherwise.

**read_hole** indicates how a partially populated region in the cache should be handled. It can be one of a number of settings:

NETFS_READ_HOLE_IGNORE - Just try to read (may return a short read).

**NETFS_READ_HOLE_CLEAR - Seek for data, clearing the part of the buffer** skipped over, then do as for IGNORE.

NETFS_READ_HOLE_FAIL - Give ENODATA if we encounter a hole.

int **fscache_begin_write_operation**(struct netfs_cache_resources *cres, struct fscache_cookie *cookie)

Begin a write operation for the netfs lib

**Parameters**

**struct netfs_cache_resources *cres**
The cache resources for the write being performed

**struct fscache_cookie *cookie**
The cookie representing the cache object

**Description**

Begin a write operation on behalf of the netfs helper library. **cres** indicates the cache resources to which the operation state should be attached; **cookie** indicates the cache object that will be accessed.

**cres->inval_counter** is set from **cookie->inval_counter** for comparison at the end of the operation. This allows invalidation during the operation to be detected by the caller.

**Return**

- 0 - Success
- **-ENOBUFS** – No caching available
- Other error code from the cache, such as -ENOMEM.

int **fscache_write**(struct netfs_cache_resources *cres, loff_t start_pos, struct iov_iter *iter, netfs_io_terminated_t term_func, void *term_func_priv)

Start a write to the cache.

**Parameters**

**struct netfs_cache_resources *cres**
The cache resources to use

**loff_t start_pos**
The beginning file offset in the cache file

**struct iov_iter *iter**
The data to write - and also the length

**netfs_io_terminated_t term_func**
The function to call upon completion

**void *term_func_priv**
The private data for **term_func**

**Description**

Start a write to the cache. **cres** indicates the cache object to write to and must be obtained by a call to fscache_begin_operation() beforehand.

The data to be written is obtained from the iterator, **iter**, and that also indicates the size of the operation. **start_pos** is the start position in the file.

Upon termination of the operation, **term_func** will be called and supplied with **term_func_priv** plus the amount of data written, if successful, or the error code otherwise.

void **fscache_clear_page_bits**(struct *address_space* *mapping, loff_t start, size_t len, bool caching)

Clear the PG_fscache bits from a set of pages

**Parameters**

**struct address_space *mapping**
The netfs inode to use as the source

**loff_t start**
The start position in **mapping**

**size_t len**
The amount of data to unlock

**bool caching**
If PG_fscache has been set

**Description**

Clear the PG_fscache flag from a sequence of pages and wake up anyone who's waiting.

void **fscache_write_to_cache**(struct fscache_cookie *cookie, struct *address_space* *mapping, loff_t start, size_t len, loff_t i_size, netfs_io_terminated_t term_func, void *term_func_priv, bool caching)

Save a write to the cache and clear PG_fscache

**Parameters**

**struct fscache_cookie *cookie**
The cookie representing the cache object

**struct address_space *mapping**
The netfs inode to use as the source

**loff_t start**
The start position in **mapping**

**size_t len**
The amount of data to write back

**loff_t i_size**
The new size of the inode

**netfs_io_terminated_t term_func**
The function to call upon completion

**void *term_func_priv**
The private data for **term_func**

**bool caching**
If PG_fscache has been set

**Description**

Helper function for a netfs to write dirty data from an inode into the cache object that's backing it.

**start** and **len** describe the range of the data. This does not need to be page-aligned, but to satisfy DIO requirements, the cache may expand it up to the page boundaries on either end. All the pages covering the range must be marked with PG_fscache.

If given, **term_func** will be called upon completion and supplied with **term_func_priv**. Note that the PG_fscache flags will have been cleared by this point, so the netfs must retain its own pin on the mapping.

void **fscache_note_page_release**(struct fscache_cookie *cookie)

Note that a netfs page got released

**Parameters**

**struct fscache_cookie *cookie**
The cookie corresponding to the file

**Description**

Note that a page that has been copied to the cache has been released. This means that future reads will need to look in the cache to see if it's there.

### 1.18.3 Cache Backend API

The FS-Cache system provides an API by which actual caches can be supplied to FS-Cache for it to then serve out to network filesystems and other interested parties. This API is used by:

```
#include <linux/fscache-cache.h>.
```

#### Overview

Interaction with the API is handled on three levels: cache, volume and data storage, and each level has its own type of cookie object:

| COOKIE | C TYPE |
|---|---|
| Cache cookie | struct fscache_cache |
| Volume cookie | struct fscache_volume |
| Data storage cookie | struct fscache_cookie |

Cookies are used to provide some filesystem data to the cache, manage state and pin the cache during access in addition to acting as reference points for the API functions. Each cookie has a debugging ID that is included in trace points to make it easier to correlate traces. Note, though, that debugging IDs are simply allocated from incrementing counters and will eventually wrap.

The cache backend and the network filesystem can both ask for cache cookies - and if they ask for one of the same name, they'll get the same cookie. Volume and data cookies, however, are created at the behest of the filesystem only.

### Cache Cookies

Caches are represented in the API by cache cookies. These are objects of type:

```
struct fscache_cache {
        void            *cache_priv;
        unsigned int    debug_id;
        char            *name;
        ...
};
```

There are a few fields that the cache backend might be interested in. The `debug_id` can be used in tracing to match lines referring to the same cache and `name` is the name the cache was registered with. The `cache_priv` member is private data provided by the cache when it is brought online. The other fields are for internal use.

### Registering a Cache

When a cache backend wants to bring a cache online, it should first register the cache name and that will get it a cache cookie. This is done with:

```
struct fscache_cache *fscache_acquire_cache(const char *name);
```

This will look up and potentially create a cache cookie. The cache cookie may have already been created by a network filesystem looking for it, in which case that cache cookie will be used. If the cache cookie is not in use by another cache, it will be moved into the preparing state, otherwise it will return busy.

If successful, the cache backend can then start setting up the cache. In the event that the initialisation fails, the cache backend should call:

```
void fscache_relinquish_cache(struct fscache_cache *cache);
```

to reset and discard the cookie.

### Bringing a Cache Online

Once the cache is set up, it can be brought online by calling:

```
int fscache_add_cache(struct fscache_cache *cache,
                      const struct fscache_cache_ops *ops,
                      void *cache_priv);
```

This stores the cache operations table pointer and cache private data into the cache cookie and moves the cache to the active state, thereby allowing accesses to take place.

### Withdrawing a Cache From Service

The cache backend can withdraw a cache from service by calling this function:

```
void fscache_withdraw_cache(struct fscache_cache *cache);
```

This moves the cache to the withdrawn state to prevent new cache- and volume-level accesses from starting and then waits for outstanding cache-level accesses to complete.

The cache must then go through the data storage objects it has and tell fscache to withdraw them, calling:

```
void fscache_withdraw_cookie(struct fscache_cookie *cookie);
```

on the cookie that each object belongs to. This schedules the specified cookie for withdrawal. This gets offloaded to a workqueue. The cache backend can wait for completion by calling:

```
void fscache_wait_for_objects(struct fscache_cache *cache);
```

Once all the cookies are withdrawn, a cache backend can withdraw all the volumes, calling:

```
void fscache_withdraw_volume(struct fscache_volume *volume);
```

to tell fscache that a volume has been withdrawn. This waits for all outstanding accesses on the volume to complete before returning.

When the cache is completely withdrawn, fscache should be notified by calling:

```
void fscache_relinquish_cache(struct fscache_cache *cache);
```

to clear fields in the cookie and discard the caller's ref on it.

### Volume Cookies

Within a cache, the data storage objects are organised into logical volumes. These are represented in the API as objects of type:

```
struct fscache_volume {
        struct fscache_cache            *cache;
        void                            *cache_priv;
        unsigned int                    debug_id;
        char                            *key;
        unsigned int                    key_hash;
        ...
        u8                              coherency_len;
        u8                              coherency[];
};
```

There are a number of fields here that are of interest to the caching backend:

- `cache` - The parent cache cookie.
- `cache_priv` - A place for the cache to stash private data.
- `debug_id` - A debugging ID for logging in tracepoints.

- key - A printable string with no '/' characters in it that represents the index key for the volume. The key is NUL-terminated and padded out to a multiple of 4 bytes.
- key_hash - A hash of the index key. This should work out the same, no matter the cpu arch and endianness.
- coherency - A piece of coherency data that should be checked when the volume is bound to in the cache.
- coherency_len - The amount of data in the coherency buffer.

### Data Storage Cookies

A volume is a logical group of data storage objects, each of which is represented to the network filesystem by a cookie. Cookies are represented in the API as objects of type:

```
struct fscache_cookie {
        struct fscache_volume           *volume;
        void                            *cache_priv;
        unsigned long                   flags;
        unsigned int                    debug_id;
        unsigned int                    inval_counter;
        loff_t                          object_size;
        u8                              advice;
        u32                             key_hash;
        u8                              key_len;
        u8                              aux_len;
        ...
};
```

The fields in the cookie that are of interest to the cache backend are:

- volume - The parent volume cookie.
- cache_priv - A place for the cache to stash private data.
- flags - A collection of bit flags, including:
  - FSCACHE_COOKIE_NO_DATA_TO_READ - There is no data available in the cache to be read as the cookie has been created or invalidated.
  - FSCACHE_COOKIE_NEEDS_UPDATE - The coherency data and/or object size has been changed and needs committing.
  - FSCACHE_COOKIE_LOCAL_WRITE - The netfs's data has been modified locally, so the cache object may be in an incoherent state with respect to the server.
  - FSCACHE_COOKIE_HAVE_DATA - The backend should set this if it successfully stores data into the cache.
  - FSCACHE_COOKIE_RETIRED - The cookie was invalidated when it was relinquished and the cached data should be discarded.
- debug_id - A debugging ID for logging in tracepoints.
- inval_counter - The number of invalidations done on the cookie.
- advice - Information about how the cookie is to be used.

- `key_hash` - A hash of the index key. This should work out the same, no matter the cpu arch and endianness.
- `key_len` - The length of the index key.
- `aux_len` - The length of the coherency data buffer.

Each cookie has an index key, which may be stored inline to the cookie or elsewhere. A pointer to this can be obtained by calling:

```
void *fscache_get_key(struct fscache_cookie *cookie);
```

The index key is a binary blob, the storage for which is padded out to a multiple of 4 bytes.

Each cookie also has a buffer for coherency data. This may also be inline or detached from the cookie and a pointer is obtained by calling:

```
void *fscache_get_aux(struct fscache_cookie *cookie);
```

### Cookie Accounting

Data storage cookies are counted and this is used to block cache withdrawal completion until all objects have been destroyed. The following functions are provided to the cache to deal with that:

```
void fscache_count_object(struct fscache_cache *cache);
void fscache_uncount_object(struct fscache_cache *cache);
void fscache_wait_for_objects(struct fscache_cache *cache);
```

The count function records the allocation of an object in a cache and the uncount function records its destruction. Warning: by the time the uncount function returns, the cache may have been destroyed.

The wait function can be used during the withdrawal procedure to wait for fscache to finish withdrawing all the objects in the cache. When it completes, there will be no remaining objects referring to the cache object or any volume objects.

### Cache Management API

The cache backend implements the cache management API by providing a table of operations that fscache can use to manage various aspects of the cache. These are held in a structure of type:

```
struct fscache_cache_ops {
        const char *name;
        ...
};
```

This contains a printable name for the cache backend driver plus a number of pointers to methods to allow fscache to request management of the cache:

- Set up a volume cookie [optional]:

```
void (*acquire_volume)(struct fscache_volume *volume);
```

This method is called when a volume cookie is being created. The caller holds a cache-level access pin to prevent the cache from going away for the duration. This method should set up the resources to access a volume in the cache and should not return until it has done so.

If successful, it can set `cache_priv` to its own data.

- Clean up volume cookie [optional]:

```
void (*free_volume)(struct fscache_volume *volume);
```

This method is called when a volume cookie is being released if `cache_priv` is set.

- Look up a cookie in the cache [mandatory]:

```
bool (*lookup_cookie)(struct fscache_cookie *cookie);
```

This method is called to look up/create the resources needed to access the data storage for a cookie. It is called from a worker thread with a volume-level access pin in the cache to prevent it from being withdrawn.

True should be returned if successful and false otherwise. If false is returned, the withdraw_cookie op (see below) will be called.

If lookup fails, but the object could still be created (e.g. it hasn't been cached before), then:

```
void fscache_cookie_lookup_negative(
        struct fscache_cookie *cookie);
```

can be called to let the network filesystem proceed and start downloading stuff whilst the cache backend gets on with the job of creating things.

If successful, `cookie->cache_priv` can be set.

- Withdraw an object without any cookie access counts held [mandatory]:

```
void (*withdraw_cookie)(struct fscache_cookie *cookie);
```

This method is called to withdraw a cookie from service. It will be called when the cookie is relinquished by the netfs, withdrawn or culled by the cache backend or closed after a period of non-use by fscache.

The caller doesn't hold any access pins, but it is called from a non-reentrant work item to manage races between the various ways withdrawal can occur.

The cookie will have the `FSCACHE_COOKIE_RETIRED` flag set on it if the associated data is to be removed from the cache.

- Change the size of a data storage object [mandatory]:

```
void (*resize_cookie)(struct netfs_cache_resources *cres,
                      loff_t new_size);
```

This method is called to inform the cache backend of a change in size of the netfs file due to local truncation. The cache backend should make all of the changes it needs to make before returning as this is done under the netfs inode mutex.

The caller holds a cookie-level access pin to prevent a race with withdrawal and the netfs must have the cookie marked in-use to prevent garbage collection or culling from removing any resources.

- Invalidate a data storage object [mandatory]:

```
bool (*invalidate_cookie)(struct fscache_cookie *cookie);
```

This is called when the network filesystem detects a third-party modification or when an O_DIRECT write is made locally. This requests that the cache backend should throw away all the data in the cache for this object and start afresh. It should return true if successful and false otherwise.

On entry, new I O/operations are blocked. Once the cache is in a position to accept I/O again, the backend should release the block by calling:

```
void fscache_resume_after_invalidation(struct fscache_cookie *cookie);
```

If the method returns false, caching will be withdrawn for this cookie.

- Prepare to make local modifications to the cache [mandatory]:

```
void (*prepare_to_write)(struct fscache_cookie *cookie);
```

This method is called when the network filesystem finds that it is going to need to modify the contents of the cache due to local writes or truncations. This gives the cache a chance to note that a cache object may be incoherent with respect to the server and may need writing back later. This may also cause the cached data to be scrapped on later rebinding if not properly committed.

- Begin an operation for the netfs lib [mandatory]:

```
bool (*begin_operation)(struct netfs_cache_resources *cres,
                        enum fscache_want_state want_state);
```

This method is called when an I/O operation is being set up (read, write or resize). The caller holds an access pin on the cookie and must have marked the cookie as in-use.

If it can, the backend should attach any resources it needs to keep around to the netfs_cache_resources object and return true.

If it can't complete the setup, it should return false.

The want_state parameter indicates the state the caller needs the cache object to be in and what it wants to do during the operation:

  - `FSCACHE_WANT_PARAMS` - The caller just wants to access cache object parameters; it doesn't need to do data I/O yet.
  - `FSCACHE_WANT_READ` - The caller wants to read data.
  - `FSCACHE_WANT_WRITE` - The caller wants to write to or resize the cache object.

Note that there won't necessarily be anything attached to the cookie's cache_priv yet if the cookie is still being created.

### Data I/O API

A cache backend provides a data I/O API by through the netfs library's `struct netfs_cache_ops` attached to a `struct netfs_cache_resources` by the `begin_operation` method described above.

See the *Network Filesystem Helper Library* for a description.

### Miscellaneous Functions

FS-Cache provides some utilities that a cache backend may make use of:

- Note occurrence of an I/O error in a cache:

```
void fscache_io_error(struct fscache_cache *cache);
```

  This tells FS-Cache that an I/O error occurred in the cache. This prevents any new I/O from being started on the cache.

  This does not actually withdraw the cache. That must be done separately.

- Note cessation of caching on a cookie due to failure:

```
void fscache_caching_failed(struct fscache_cookie *cookie);
```

  This notes that a the caching that was being done on a cookie failed in some way, for instance the backing storage failed to be created or invalidation failed and that no further I/O operations should take place on it until the cache is reset.

- Count I/O requests:

```
void fscache_count_read(void);
void fscache_count_write(void);
```

  These record reads and writes from/to the cache. The numbers are displayed in /proc/fs/fscache/stats.

- Count out-of-space errors:

```
void fscache_count_no_write_space(void);
void fscache_count_no_create_space(void);
```

  These record ENOSPC errors in the cache, divided into failures of data writes and failures of filesystem object creations (e.g. mkdir).

- Count objects culled:

```
void fscache_count_culled(void);
```

  This records the culling of an object.

- Get the cookie from a set of cache resources:

```
struct fscache_cookie *fscache_cres_cookie(struct netfs_cache_resources␣
↪*cres)
```

Pull a pointer to the cookie from the cache resources. This may return a NULL cookie if no cookie was set.

### API Function Reference

enum *fscache_cookie_state* **fscache_cookie_state**(struct fscache_cookie *cookie)

Read the state of a cookie

**Parameters**

**struct fscache_cookie *cookie**
The cookie to query

**Description**

Get the state of a cookie, imposing an ordering between the cookie contents and the state value. Paired with fscache_set_cookie_state().

void ***fscache_get_key**(struct fscache_cookie *cookie)

Get a pointer to the cookie key

**Parameters**

**struct fscache_cookie *cookie**
The cookie to query

**Description**

Return a pointer to the where a cookie's key is stored.

void **fscache_count_object**(struct fscache_cache *cache)

Tell fscache that an object has been added

**Parameters**

**struct fscache_cache *cache**
The cache to account to

**Description**

Tell fscache that an object has been added to the cache. This prevents the cache from tearing down the cache structure until the object is uncounted.

void **fscache_uncount_object**(struct fscache_cache *cache)

Tell fscache that an object has been removed

**Parameters**

**struct fscache_cache *cache**
The cache to account to

**Description**

Tell fscache that an object has been removed from the cache and will no longer be accessed. After this point, the cache cookie may be destroyed.

void **fscache_wait_for_objects**(struct fscache_cache *cache)

Wait for all objects to be withdrawn

**Parameters**

**struct fscache_cache *cache**
The cache to query

**Description**

Wait for all extant objects in a cache to finish being withdrawn and go away.

### 1.18.4 Cache on Already Mounted Filesystem

#### Overview

CacheFiles is a caching backend that's meant to use as a cache a directory on an already mounted filesystem of a local type (such as Ext3).

CacheFiles uses a userspace daemon to do some of the cache management - such as reaping stale nodes and culling. This is called cachefilesd and lives in /sbin.

The filesystem and data integrity of the cache are only as good as those of the filesystem providing the backing services. Note that CacheFiles does not attempt to journal anything since the journalling interfaces of the various filesystems are very specific in nature.

CacheFiles creates a misc character device - "/dev/cachefiles" - that is used to communication with the daemon. Only one thing may have this open at once, and while it is open, a cache is at least partially in existence. The daemon opens this and sends commands down it to control the cache.

CacheFiles is currently limited to a single cache.

CacheFiles attempts to maintain at least a certain percentage of free space on the filesystem, shrinking the cache by culling the objects it contains to make space if necessary - see the "Cache Culling" section. This means it can be placed on the same medium as a live set of data, and will expand to make use of spare space and automatically contract when the set of data requires more space.

#### Requirements

The use of CacheFiles and its daemon requires the following features to be available in the system and in the cache filesystem:

- dnotify.
- extended attributes (xattrs).
- openat() and friends.
- *bmap()* support on files in the filesystem (FIBMAP ioctl).
- The use of *bmap()* to detect a partial page at the end of the file.

It is strongly recommended that the "dir_index" option is enabled on Ext3 filesystems being used as a cache.

### Configuration

The cache is configured by a script in /etc/cachefilesd.conf. These commands set up cache ready for use. The following script commands are available:

**brun <N>%, bcull <N>%, bstop <N>%, frun <N>%, fcull <N>%, fstop <N>%**
Configure the culling limits. Optional. See the section on culling The defaults are 7% (run), 5% (cull) and 1% (stop) respectively.

The commands beginning with a 'b' are file space (block) limits, those beginning with an 'f' are file count limits.

**dir <path>**
Specify the directory containing the root of the cache. Mandatory.

**tag <name>**
Specify a tag to FS-Cache to use in distinguishing multiple caches. Optional. The default is "CacheFiles".

**debug <mask>**
Specify a numeric bitmask to control debugging in the kernel module. Optional. The default is zero (all off). The following values can be OR'd into the mask to collect various information:

| | |
|---|---|
| 1 | Turn on trace of function entry (_enter() macros) |
| 2 | Turn on trace of function exit (_leave() macros) |
| 4 | Turn on trace of internal debug points (_debug()) |

This mask can also be set through sysfs, eg:

```
echo 5 >/sys/modules/cachefiles/parameters/debug
```

### Starting the Cache

The cache is started by running the daemon. The daemon opens the cache device, configures the cache and tells it to begin caching. At that point the cache binds to fscache and the cache becomes live.

The daemon is run as follows:

```
/sbin/cachefilesd [-d]* [-s] [-n] [-f <configfile>]
```

The flags are:

**-d**
Increase the debugging level. This can be specified multiple times and is cumulative with itself.

**-s**
Send messages to stderr instead of syslog.

**-n**
Don't daemonise and go into background.

**-f <configfile>**
Use an alternative configuration file rather than the default one.

### Things to Avoid

Do not mount other things within the cache as this will cause problems. The kernel module contains its own very cut-down path walking facility that ignores mountpoints, but the daemon can't avoid them.

Do not create, rename or unlink files and directories in the cache while the cache is active, as this may cause the state to become uncertain.

Renaming files in the cache might make objects appear to be other objects (the filename is part of the lookup key).

Do not change or remove the extended attributes attached to cache files by the cache as this will cause the cache state management to get confused.

Do not create files or directories in the cache, lest the cache get confused or serve incorrect data.

Do not chmod files in the cache. The module creates things with minimal permissions to prevent random users being able to access them directly.

### Cache Culling

The cache may need culling occasionally to make space. This involves discarding objects from the cache that have been used less recently than anything else. Culling is based on the access time of data objects. Empty directories are culled if not in use.

Cache culling is done on the basis of the percentage of blocks and the percentage of files available in the underlying filesystem. There are six "limits":

**brun, frun**
If the amount of free space and the number of available files in the cache rises above both these limits, then culling is turned off.

**bcull, fcull**
If the amount of available space or the number of available files in the cache falls below either of these limits, then culling is started.

**bstop, fstop**
If the amount of available space or the number of available files in the cache falls below either of these limits, then no further allocation of disk space or files is permitted until culling has raised things above these limits again.

These must be configured thusly:

```
0 <= bstop < bcull < brun < 100
0 <= fstop < fcull < frun < 100
```

Note that these are percentages of available space and available files, and do _not_ appear as 100 minus the percentage displayed by the "df" program.

The userspace daemon scans the cache to build up a table of cullable objects. These are then culled in least recently used order. A new scan of the cache is started as soon as space is made

in the table. Objects will be skipped if their atimes have changed or if the kernel module says it is still using them.

### Cache Structure

The CacheFiles module will create two directories in the directory it was given:

- cache/
- graveyard/

The active cache objects all reside in the first directory. The CacheFiles kernel module moves any retired or culled objects that it can't simply unlink to the graveyard from which the daemon will actually delete them.

The daemon uses dnotify to monitor the graveyard directory, and will delete anything that appears therein.

The module represents index objects as directories with the filename "I..." or "J...". Note that the "cache/" directory is itself a special index.

Data objects are represented as files if they have no children, or directories if they do. Their filenames all begin "D..." or "E...". If represented as a directory, data objects will have a file in the directory called "data" that actually holds the data.

Special objects are similar to data objects, except their filenames begin "S..." or "T...".

If an object has children, then it will be represented as a directory. Immediately in the representative directory are a collection of directories named for hash values of the child object keys with an '@' prepended. Into this directory, if possible, will be placed the representations of the child objects:

```
 /INDEX    /INDEX     /INDEX                                  /DATA FILES
/=========/==========/================================/================
cache/@4a/I03nfs/@30/Ji000000000000000--fHg8hi8400
cache/@4a/I03nfs/@30/Ji000000000000000--fHg8hi8400/@75/Es0g000w...DB1ry
cache/@4a/I03nfs/@30/Ji000000000000000--fHg8hi8400/@75/Es0g000w...N22ry
cache/@4a/I03nfs/@30/Ji000000000000000--fHg8hi8400/@75/Es0g000w...FP1ry
```

If the key is so long that it exceeds NAME_MAX with the decorations added on to it, then it will be cut into pieces, the first few of which will be used to make a nest of directories, and the last one of which will be the objects inside the last directory. The names of the intermediate directories will have '+' prepended:

```
J1223/@23/+xy...z/+kl...m/Epqr
```

Note that keys are raw data, and not only may they exceed NAME_MAX in size, they may also contain things like '/' and NUL characters, and so they may not be suitable for turning directly into a filename.

To handle this, CacheFiles will use a suitably printable filename directly and "base-64" encode ones that aren't directly suitable. The two versions of object filenames indicate the encoding:

| OBJECT TYPE | PRINTABLE | ENCODED |
|---|---|---|
| Index | ”I...” | ”J...” |
| Data | ”D...” | ”E...” |
| Special | ”S...” | ”T...” |

Intermediate directories are always ”@” or ”+” as appropriate.

Each object in the cache has an extended attribute label that holds the object type ID (required to distinguish special objects) and the auxiliary data from the netfs. The latter is used to detect stale objects in the cache and update or retire them.

Note that CacheFiles will erase from the cache any file it doesn't recognise or any file of an incorrect type (such as a FIFO file or a device file).

### Security Model and SELinux

CacheFiles is implemented to deal properly with the LSM security features of the Linux kernel and the SELinux facility.

One of the problems that CacheFiles faces is that it is generally acting on behalf of a process, and running in that process's context, and that includes a security context that is not appropriate for accessing the cache - either because the files in the cache are inaccessible to that process, or because if the process creates a file in the cache, that file may be inaccessible to other processes.

The way CacheFiles works is to temporarily change the security context (fsuid, fsgid and actor security label) that the process acts as - without changing the security context of the process when it the target of an operation performed by some other process (so signalling and suchlike still work correctly).

When the CacheFiles module is asked to bind to its cache, it:

(1) Finds the security label attached to the root cache directory and uses that as the security label with which it will create files. By default, this is:

```
cachefiles_var_t
```

(2) Finds the security label of the process which issued the bind request (presumed to be the cachefilesd daemon), which by default will be:

```
cachefilesd_t
```

and asks LSM to supply a security ID as which it should act given the daemon's label. By default, this will be:

```
cachefiles_kernel_t
```

SELinux transitions the daemon's security ID to the module's security ID based on a rule of this form in the policy:

```
type_transition <daemon's-ID> kernel_t : process <module's-ID>;
```

For instance:

```
type_transition cachefilesd_t kernel_t : process cachefiles_kernel_t;
```

The module's security ID gives it permission to create, move and remove files and directories in the cache, to find and access directories and files in the cache, to set and access extended attributes on cache objects, and to read and write files in the cache.

The daemon's security ID gives it only a very restricted set of permissions: it may scan directories, stat files and erase files and directories. It may not read or write files in the cache, and so it is precluded from accessing the data cached therein; nor is it permitted to create new files in the cache.

There are policy source files available in:

https://people.redhat.com/~dhowells/fscache/cachefilesd-0.8.tar.bz2

and later versions. In that tarball, see the files:

```
cachefilesd.te
cachefilesd.fc
cachefilesd.if
```

They are built and installed directly by the RPM.

If a non-RPM based system is being used, then copy the above files to their own directory and run:

```
make -f /usr/share/selinux/devel/Makefile
semodule -i cachefilesd.pp
```

You will need checkpolicy and selinux-policy-devel installed prior to the build.

By default, the cache is located in /var/fscache, but if it is desirable that it should be elsewhere, than either the above policy files must be altered, or an auxiliary policy must be installed to label the alternate location of the cache.

For instructions on how to add an auxiliary policy to enable the cache to be located elsewhere when SELinux is in enforcing mode, please see:

```
/usr/share/doc/cachefilesd-*/move-cache.txt
```

When the cachefilesd rpm is installed; alternatively, the document can be found in the sources.

### A Note on Security

CacheFiles makes use of the split security in the task_struct. It allocates its own task_security structure, and redirects current->cred to point to it when it acts on behalf of another process, in that process's context.

The reason it does this is that it calls *vfs_mkdir()* and suchlike rather than bypassing security and calling inode ops directly. Therefore the VFS and LSM may deny the CacheFiles access to the cache data because under some circumstances the caching code is running in the security context of whatever process issued the original syscall on the netfs.

Furthermore, should CacheFiles create a file or directory, the security parameters with that object is created (UID, GID, security label) would be derived from that process that issued the

system call, thus potentially preventing other processes from accessing the cache - including CacheFiles's cache management daemon (cachefilesd).

What is required is to temporarily override the security of the process that issued the system call. We can't, however, just do an in-place change of the security data as that affects the process as an object, not just as a subject. This means it may lose signals or ptrace events for example, and affects what the process looks like in /proc.

So CacheFiles makes use of a logical split in the security between the objective security (task->real_cred) and the subjective security (task->cred). The objective security holds the intrinsic security properties of a process and is never overridden. This is what appears in /proc, and is what is used when a process is the target of an operation by some other process (SIGKILL for example).

The subjective security holds the active security properties of a process, and may be overridden. This is not seen externally, and is used when a process acts upon another object, for example SIGKILLing another process or opening a file.

LSM hooks exist that allow SELinux (or Smack or whatever) to reject a request for CacheFiles to run in a context of a specific security label, or to create files and directories with another security label.

### Statistical Information

If FS-Cache is compiled with the following option enabled:

```
CONFIG_CACHEFILES_HISTOGRAM=y
```

then it will gather certain statistics and display them through a proc file.

/proc/fs/cachefiles/histogram

```
cat /proc/fs/cachefiles/histogram
JIFS  SECS  LOOKUPS   MKDIRS    CREATES
===== ===== ========= ========= =========
```

This shows the breakdown of the number of times each amount of time between 0 jiffies and HZ-1 jiffies a variety of tasks took to run. The columns are as follows:

| COLUMN | TIME MEASUREMENT |
|---|---|
| LOOKUPS | Length of time to perform a lookup on the backing fs |
| MKDIRS | Length of time to perform a mkdir on the backing fs |
| CREATES | Length of time to perform a create on the backing fs |

Each row shows the number of events that took a particular range of times. Each step is 1 jiffy in size. The JIFS column indicates the particular jiffy range covered, and the SECS field the equivalent number of seconds.

### Debugging

If CONFIG_CACHEFILES_DEBUG is enabled, the CacheFiles facility can have runtime debugging enabled by adjusting the value in:

```
/sys/module/cachefiles/parameters/debug
```

This is a bitmask of debugging streams to enable:

| BIT | VALUE | STREAM | POINT |
|---|---|---|---|
| 0 | 1 | General | Function entry trace |
| 1 | 2 | | Function exit trace |
| 2 | 4 | | General |

The appropriate set of values should be OR'd together and the result written to the control file. For example:

```
echo $((1|4|8)) >/sys/module/cachefiles/parameters/debug
```

will turn on all function entry debugging.

### On-demand Read

When working in its original mode, CacheFiles serves as a local cache for a remote networking fs - while in on-demand read mode, CacheFiles can boost the scenario where on-demand read semantics are needed, e.g. container image distribution.

The essential difference between these two modes is seen when a cache miss occurs: In the original mode, the netfs will fetch the data from the remote server and then write it to the cache file; in on-demand read mode, fetching the data and writing it into the cache is delegated to a user daemon.

`CONFIG_CACHEFILES_ONDEMAND` should be enabled to support on-demand read mode.

### Protocol Communication

The on-demand read mode uses a simple protocol for communication between kernel and user daemon. The protocol can be modeled as:

```
kernel --[request]--> user daemon --[reply]--> kernel
```

CacheFiles will send requests to the user daemon when needed. The user daemon should poll the devnode ('/dev/cachefiles') to check if there's a pending request to be processed. A POLLIN event will be returned when there's a pending request.

The user daemon then reads the devnode to fetch a request to process. It should be noted that each read only gets one request. When it has finished processing the request, the user daemon should write the reply to the devnode.

Each request starts with a message header of the form:

```
struct cachefiles_msg {
        __u32 msg_id;
        __u32 opcode;
        __u32 len;
        __u32 object_id;
        __u8  data[];
};
```

where:

- `msg_id` is a unique ID identifying this request among all pending requests.
- `opcode` indicates the type of this request.
- `object_id` is a unique ID identifying the cache file operated on.
- `data` indicates the payload of this request.
- `len` indicates the whole length of this request, including the header and following type-specific payload.

### Turning on On-demand Mode

An optional parameter becomes available to the ”bind” command:

```
bind [ondemand]
```

When the ”bind” command is given no argument, it defaults to the original mode. When it is given the ”ondemand” argument, i.e. ”bind ondemand”, on-demand read mode will be enabled.

### The OPEN Request

When the netfs opens a cache file for the first time, a request with the CACHEFILES_OP_OPEN opcode, a.k.a an OPEN request will be sent to the user daemon. The payload format is of the form:

```
struct cachefiles_open {
        __u32 volume_key_size;
        __u32 cookie_key_size;
        __u32 fd;
        __u32 flags;
        __u8  data[];
};
```

where:

- `data` contains the volume_key followed directly by the cookie_key. The volume key is a NUL-terminated string; the cookie key is binary data.
- `volume_key_size` indicates the size of the volume key in bytes.
- `cookie_key_size` indicates the size of the cookie key in bytes.

- `fd` indicates an anonymous fd referring to the cache file, through which the user daemon can perform write/llseek file operations on the cache file.

The user daemon can use the given (volume_key, cookie_key) pair to distinguish the requested cache file. With the given anonymous fd, the user daemon can fetch the data and write it to the cache file in the background, even when kernel has not triggered a cache miss yet.

Be noted that each cache file has a unique object_id, while it may have multiple anonymous fds. The user daemon may duplicate anonymous fds from the initial anonymous fd indicated by the @fd field through dup(). Thus each object_id can be mapped to multiple anonymous fds, while the usr daemon itself needs to maintain the mapping.

When implementing a user daemon, please be careful of RLIMIT_NOFILE, `/proc/sys/fs/nr_open` and `/proc/sys/fs/file-max`. Typically these needn't be huge since they're related to the number of open device blobs rather than open files of each individual filesystem.

The user daemon should reply the OPEN request by issuing a ”copen” (complete open) command on the devnode:

```
copen <msg_id>,<cache_size>
```

where:

- `msg_id` must match the msg_id field of the OPEN request.
- When >= 0, `cache_size` indicates the size of the cache file; when < 0, `cache_size` indicates any error code encountered by the user daemon.

### The CLOSE Request

When a cookie withdrawn, a CLOSE request (opcode CACHEFILES_OP_CLOSE) will be sent to the user daemon. This tells the user daemon to close all anonymous fds associated with the given object_id. The CLOSE request has no extra payload, and shouldn't be replied.

### The READ Request

When a cache miss is encountered in on-demand read mode, CacheFiles will send a READ request (opcode CACHEFILES_OP_READ) to the user daemon. This tells the user daemon to fetch the contents of the requested file range. The payload is of the form:

```
struct cachefiles_read {
        __u64 off;
        __u64 len;
};
```

where:

- `off` indicates the starting offset of the requested file range.
- `len` indicates the length of the requested file range.

When it receives a READ request, the user daemon should fetch the requested data and write it to the cache file identified by object_id.

When it has finished processing the READ request, the user daemon should reply by using the CACHEFILES_IOC_READ_COMPLETE ioctl on one of the anonymous fds associated with the object_id given in the READ request. The ioctl is of the form:

```
ioctl(fd, CACHEFILES_IOC_READ_COMPLETE, msg_id);
```

where:

- `fd` is one of the anonymous fds associated with the object_id given.
- `msg_id` must match the msg_id field of the READ request.

## 1.19 Changes since 2.5.0:

---

**recommended**

New helpers: sb_bread(), sb_getblk(), sb_find_get_block(), set_bh(), sb_set_blocksize() and sb_min_blocksize().

Use them.

(sb_find_get_block() replaces 2.4's get_hash_table())

---

**recommended**

New methods: ->alloc_inode() and ->destroy_inode().

Remove inode->u.foo_inode_i

Declare:

```
struct foo_inode_info {
        /* fs-private stuff */
        struct inode vfs_inode;
};
static inline struct foo_inode_info *FOO_I(struct inode *inode)
{
        return list_entry(inode, struct foo_inode_info, vfs_inode);
}
```

Use FOO_I(inode) instead of &inode->u.foo_inode_i;

Add foo_alloc_inode() and foo_destroy_inode() - the former should allocate foo_inode_info and return the address of ->vfs_inode, the latter should free FOO_I(inode) (see in-tree filesystems for examples).

Make them ->alloc_inode and ->destroy_inode in your super_operations.

Keep in mind that now you need explicit initialization of private data typically between calling *iget_locked()* and unlocking the inode.

At some point that will become mandatory.

**mandatory**

The foo_inode_info should always be allocated through alloc_inode_sb() rather than kmem_cache_alloc() or kmalloc() related to set up the inode reclaim context correctly.

---

**mandatory**

Change of file_system_type method (->read_super to ->get_sb)

->read_super() is no more. Ditto for DECLARE_FSTYPE and DECLARE_FSTYPE_DEV.

Turn your foo_read_super() into a function that would return 0 in case of success and negative number in case of error (-EINVAL unless you have more informative error value to report). Call it foo_fill_super(). Now declare:

```
int foo_get_sb(struct file_system_type *fs_type,
      int flags, const char *dev_name, void *data, struct vfsmount *mnt)
{
      return get_sb_bdev(fs_type, flags, dev_name, data, foo_fill_super,
                         mnt);
}
```

(or similar with s/bdev/nodev/ or s/bdev/single/, depending on the kind of filesystem).

Replace DECLARE_FSTYPE... with explicit initializer and have ->get_sb set as foo_get_sb.

---

**mandatory**

Locking change: ->s_vfs_rename_sem is taken only by cross-directory renames. Most likely there is no need to change anything, but if you relied on global exclusion between renames for some internal purpose - you need to change your internal locking. Otherwise exclusion warranties remain the same (i.e. parents and victim are locked, etc.).

---

**informational**

Now we have the exclusion between ->lookup() and directory removal (by ->rmdir() and ->rename()). If you used to need that exclusion and do it by internal locking (most of filesystems couldn't care less) - you can relax your locking.

---

**mandatory**

->lookup(), ->truncate(), ->create(), ->unlink(), ->mknod(), ->mkdir(), ->rmdir(), ->link(), ->lseek(), ->symlink(), ->rename() and ->readdir() are called without BKL now. Grab it on entry, drop upon return - that will guarantee the same locking you used to have. If your method or its parts do not need BKL - better yet, now you can shift lock_kernel() and unlock_kernel() so that they would protect exactly what needs to be protected.

---

**mandatory**

BKL is also moved from around sb operations. BKL should have been shifted into individual fs sb_op functions. If you don't need it, remove it.

---

**informational**

check for ->link() target not being a directory is done by callers. Feel free to drop it...

---

**informational**

->link() callers hold ->i_mutex on the object we are linking to. Some of your problems might be over...

---

**mandatory**

new file_system_type method - kill_sb(superblock). If you are converting an existing filesystem, set it according to ->fs_flags:

```
FS_REQUIRES_DEV         -       kill_block_super
FS_LITTER               -       kill_litter_super
neither                 -       kill_anon_super
```

FS_LITTER is gone - just remove it from fs_flags.

---

**mandatory**

FS_SINGLE is gone (actually, that had happened back when ->get_sb() went in - and hadn't been documented ;-/). Just remove it from fs_flags (and see ->get_sb() entry for other actions).

---

**mandatory**

->setattr() is called without BKL now. Caller _always_ holds ->i_mutex, so watch for ->i_mutex-grabbing code that might be used by your ->setattr(). Callers of *notify_change()* need ->i_mutex now.

---

**recommended**

New super_block field `struct export_operations *s_export_op` for explicit support for exporting, e.g. via NFS. The structure is fully documented at its declaration in include/linux/fs.h, and in *Making Filesystems Exportable*.

Briefly it allows for the definition of decode_fh and encode_fh operations to encode and decode filehandles, and allows the filesystem to use a standard helper function for decode_fh, and provide file-system specific support for this helper, particularly get_parent.

It is planned that this will be required for exporting once the code settles down a bit.

**mandatory**

s_export_op is now required for exporting a filesystem. isofs, ext2, ext3, reiserfs, fat can be used as examples of very different filesystems.

---

**mandatory**

iget4() and the read_inode2 callback have been superseded by *iget5_locked()* which has the following prototype:

```
struct inode *iget5_locked(struct super_block *sb, unsigned long ino,
                           int (*test)(struct inode *, void *),
                           int (*set)(struct inode *, void *),
                           void *data);
```

'test' is an additional function that can be used when the inode number is not sufficient to identify the actual file object. 'set' should be a non-blocking function that initializes those parts of a newly created inode to allow the test function to succeed. 'data' is passed as an opaque value to both test and set functions.

When the inode has been created by *iget5_locked()*, it will be returned with the I_NEW flag set and will still be locked. The filesystem then needs to finalize the initialization. Once the inode is initialized it must be unlocked by calling *unlock_new_inode()*.

The filesystem is responsible for setting (and possibly testing) i_ino when appropriate. There is also a simpler iget_locked function that just takes the superblock and inode number as arguments and does the test and set for you.

e.g.:

```
inode = iget_locked(sb, ino);
if (inode->i_state & I_NEW) {
        err = read_inode_from_disk(inode);
        if (err < 0) {
                iget_failed(inode);
                return err;
        }
        unlock_new_inode(inode);
}
```

Note that if the process of setting up a new inode fails, then *iget_failed()* should be called on the inode to render it dead, and an appropriate error should be passed back to the caller.

---

**recommended**

->getattr() finally getting used. See instances in nfs, minix, etc.

---

**mandatory**

->revalidate() is gone. If your filesystem had it - provide ->getattr() and let it call whatever you had as ->revlidate() + (for symlinks that had ->revalidate()) add calls in ->follow_link()/->readlink().

---

**mandatory**

->d_parent changes are not protected by BKL anymore. Read access is safe if at least one of the following is true:

- filesystem has no cross-directory rename()

- we know that parent had been locked (e.g. we are looking at ->d_parent of ->lookup() argument).
- we are called from ->rename().
- the child's ->d_lock is held

Audit your code and add locking if needed. Notice that any place that is not protected by the conditions above is risky even in the old tree - you had been relying on BKL and that's prone to screwups. Old tree had quite a few holes of that kind - unprotected access to ->d_parent leading to anything from oops to silent memory corruption.

---

**mandatory**

FS_NOMOUNT is gone. If you use it - just set SB_NOUSER in flags (see rootfs for one kind of solution and bdev/socket/pipe for another).

---

**recommended**

Use bdev_read_only(bdev) instead of is_read_only(kdev). The latter is still alive, but only because of the mess in drivers/s390/block/dasd.c. As soon as it gets fixed is_read_only() will die.

---

**mandatory**

->permission() is called without BKL now. Grab it on entry, drop upon return - that will guarantee the same locking you used to have. If your method or its parts do not need BKL - better yet, now you can shift lock_kernel() and unlock_kernel() so that they would protect exactly what needs to be protected.

---

**mandatory**

->statfs() is now called without BKL held. BKL should have been shifted into individual fs sb_op functions where it's not clear that it's safe to remove it. If you don't need it, remove it.

---

**mandatory**

is_read_only() is gone; use bdev_read_only() instead.

---

**mandatory**

destroy_buffers() is gone; use invalidate_bdev().

---

**mandatory**

fsync_dev() is gone; use fsync_bdev(). NOTE: lvm breakage is deliberate; as soon as struct block_device * is propagated in a reasonable way by that code fixing will become trivial; until then nothing can be done.

**mandatory**

block truncatation on error exit from ->write_begin, and ->direct_IO moved from generic methods (block_write_begin, cont_write_begin, nobh_write_begin, blockdev_direct_IO*) to callers. Take a look at ext2_write_failed and callers for an example.

**mandatory**

->truncate is gone. The whole truncate sequence needs to be implemented in ->setattr, which is now mandatory for filesystems implementing on-disk size changes. Start with a copy of the old inode_setattr and vmtruncate, and the reorder the vmtruncate + foofs_vmtruncate sequence to be in order of zeroing blocks using block_truncate_page or similar helpers, size update and on finally on-disk truncation which should not fail. setattr_prepare (which used to be inode_change_ok) now includes the size checks for ATTR_SIZE and must be called in the beginning of ->setattr unconditionally.

**mandatory**

->clear_inode() and ->delete_inode() are gone; ->evict_inode() should be used instead. It gets called whenever the inode is evicted, whether it has remaining links or not. Caller does *not* evict the pagecache or inode-associated metadata buffers; the method has to use truncate_inode_pages_final() to get rid of those. Caller makes sure async writeback cannot be running for the inode while (or after) ->evict_inode() is called.

->drop_inode() returns int now; it's called on final *iput()* with inode->i_lock held and it returns true if filesystems wants the inode to be dropped. As before, generic_drop_inode() is still the default and it's been updated appropriately. generic_delete_inode() is also alive and it consists simply of return 1. Note that all actual eviction work is done by caller after ->drop_inode() returns.

As before, clear_inode() must be called exactly once on each call of ->evict_inode() (as it used to be for each call of ->delete_inode()). Unlike before, if you are using inode-associated metadata buffers (i.e. mark_buffer_dirty_inode()), it's your responsibility to call invalidate_inode_buffers() before clear_inode().

NOTE: checking i_nlink in the beginning of ->write_inode() and bailing out if it's zero is not *and never had been* enough. Final unlink() and *iput()* may happen while the inode is in the middle of ->write_inode(); e.g. if you blindly free the on-disk inode, you may end up doing that while ->write_inode() is writing to it.

---

**mandatory**

*.d_delete()* now only advises the dcache as to whether or not to cache unreferenced dentries, and is now only called when the dentry refcount goes to 0. Even on 0 refcount transition, it must be able to tolerate being called 0, 1, or more times (eg. constant, idempotent).

---

**mandatory**

.d_compare() calling convention and locking rules are significantly changed. Read updated documentation in *Overview of the Linux Virtual File System* (and look at examples of other filesystems) for guidance.

---

**mandatory**

.d_hash() calling convention and locking rules are significantly changed. Read updated documentation in *Overview of the Linux Virtual File System* (and look at examples of other filesystems) for guidance.

---

**mandatory**

dcache_lock is gone, replaced by fine grained locks. See fs/dcache.c for details of what locks to replace dcache_lock with in order to protect particular things. Most of the time, a filesystem only needs ->d_lock, which protects *all* the dcache state of a given dentry.

---

**mandatory**

Filesystems must RCU-free their inodes, if they can have been accessed via rcu-walk path walk (basically, if the file can have had a path name in the vfs namespace).

Even though i_dentry and i_rcu share storage in a union, we will initialize the former in `inode_init_always()`, so just leave it alone in the callback. It used to be necessary to clean it there, but not anymore (starting at 3.2).

---

**recommended**

vfs now tries to do path walking in "rcu-walk mode", which avoids atomic operations and scalability hazards on dentries and inodes (see *Pathname lookup*). d_hash and d_compare changes (above) are examples of the changes required to support this. For more complex filesystem callbacks, the vfs drops out of rcu-walk mode before the fs call, so no changes are required to the filesystem. However, this is costly and loses the benefits of rcu-walk mode. We will begin to add filesystem callbacks that are rcu-walk aware, shown below. Filesystems should take advantage of this where possible.

---

**mandatory**

d_revalidate is a callback that is made on every path element (if the filesystem provides it), which requires dropping out of rcu-walk mode. This may now be called in rcu-walk mode (nd->flags & LOOKUP_RCU). -ECHILD should be returned if the filesystem cannot handle rcu-walk. See *Overview of the Linux Virtual File System* for more details.

permission is an inode permission check that is called on many or all directory inodes on the way down a path walk (to check for exec permission). It must now be rcu-walk aware (mask & MAY_NOT_BLOCK). See *Overview of the Linux Virtual File System* for more details.

---

**mandatory**

In ->fallocate() you must check the mode option passed in. If your filesystem does not support hole punching (deallocating space in the middle of a file) you must return -EOPNOTSUPP if FALLOC_FL_PUNCH_HOLE is set in mode. Currently you can only have FALLOC_FL_PUNCH_HOLE with FALLOC_FL_KEEP_SIZE set, so the i_size should not change when hole punching, even when puching the end of a file off.

---

**mandatory**

->get_sb() is gone. Switch to use of ->mount(). Typically it's just a matter of switching from calling `get_sb_...` to `mount_...` and changing the function type. If you were doing it manually, just switch from setting ->mnt_root to some pointer to returning that pointer. On errors return ERR_PTR(...).

---

**mandatory**

->permission() and *generic_permission()*have lost flags argument; instead of passing IPERM_FLAG_RCU we add MAY_NOT_BLOCK into mask.

*generic_permission()* has also lost the check_acl argument; ACL checking has been taken to VFS and filesystems need to provide a non-NULL ->i_op->get_inode_acl to read an ACL from disk.

---

**mandatory**

If you implement your own ->llseek() you must handle SEEK_HOLE and SEEK_DATA. You can handle this by returning -EINVAL, but it would be nicer to support it in some way. The generic handler assumes that the entire file is data and there is a virtual hole at the end of the file. So if the provided offset is less than i_size and SEEK_DATA is specified, return the same offset. If the above is true for the offset and you are given SEEK_HOLE, return the end of the file. If the offset is i_size or greater return -ENXIO in either case.

**mandatory**

If you have your own ->fsync() you must make sure to call filemap_write_and_wait_range() so that all dirty pages are synced out properly. You must also keep in mind that ->fsync() is not called with i_mutex held anymore, so if you require i_mutex locking you must make sure to take it and release it yourself.

---

**mandatory**

d_alloc_root() is gone, along with a lot of bugs caused by code misusing it. Replacement: d_make_root(inode). On success d_make_root(inode) allocates and returns a new dentry instantiated with the passed in inode. On failure NULL is returned and the passed in inode is dropped so the reference to inode is consumed in all cases and failure handling need not do any cleanup for the inode. If d_make_root(inode) is passed a NULL inode it returns NULL and also requires no further error handling. Typical usage is:

```
inode = foofs_new_inode(....);
s->s_root = d_make_root(inode);
if (!s->s_root)
        /* Nothing needed for the inode cleanup */
        return -ENOMEM;
...
```

---

**mandatory**

The witch is dead! Well, 2/3 of it, anyway. ->d_revalidate() and ->lookup() do *not* take struct nameidata anymore; just the flags.

---

**mandatory**

->create() doesn't take `struct nameidata *`; unlike the previous two, it gets "is it an O_EXCL or equivalent?" boolean argument. Note that local filesystems can ignore this argument - they are guaranteed that the object doesn't exist. It's remote/distributed ones that might care...

---

**mandatory**

FS_REVAL_DOT is gone; if you used to have it, add ->d_weak_revalidate() in your dentry operations instead.

---

**mandatory**

vfs_readdir() is gone; switch to iterate_dir() instead

---

**mandatory**

->readdir() is gone now; switch to ->iterate_shared()

**mandatory**

vfs_follow_link has been removed. Filesystems must use nd_set_link from ->follow_link for normal symlinks, or nd_jump_link for magic /proc/<pid> style links.

---

**mandatory**

*iget5_locked()*/*ilookup5()*/*ilookup5_nowait()* test() callback used to be called with both ->i_lock and inode_hash_lock held; the former is *not* taken anymore, so verify that your callbacks do not rely on it (none of the in-tree instances did). inode_hash_lock is still held, of course, so they are still serialized wrt removal from inode hash, as well as wrt set() callback of *iget5_locked()*.

---

**mandatory**

d_materialise_unique() is gone; *d_splice_alias()* does everything you need now. Remember that they have opposite orders of arguments ;-/

---

**mandatory**

f_dentry is gone; use f_path.dentry, or, better yet, see if you can avoid it entirely.

---

**mandatory**

never call ->read() and ->write() directly; use __vfs_{read,write} or wrappers; instead of checking for ->write or ->read being NULL, look for FMODE_CAN_{WRITE,READ} in file->f_mode.

---

**mandatory**

do _not_ use new_sync_{read,write} for ->read/->write; leave it NULL instead.

---

**mandatory**

->aio_read/->aio_write are gone. Use ->read_iter/->write_iter.

---

**recommended**

for embedded ("fast") symlinks just set inode->i_link to wherever the symlink body is and use simple_follow_link() as ->follow_link().

---

**mandatory**

calling conventions for ->follow_link() have changed. Instead of returning cookie and using nd_set_link() to store the body to traverse, we return the body to traverse and store the cookie using explicit void ** argument. nameidata isn't passed at all - nd_jump_link() doesn't need it and nd_[gs]et_link() is gone.

---

**mandatory**

calling conventions for ->put_link() have changed. It gets inode instead of dentry, it does not get nameidata at all and it gets called only when cookie is non-NULL. Note that link body isn't available anymore, so if you need it, store it as cookie.

---

**mandatory**

any symlink that might use page_follow_link_light/page_put_link() must have inode_nohighmem(inode) called before anything might start playing with its pagecache. No highmem pages should end up in the pagecache of such symlinks. That includes any preseeding that might be done during symlink creation. page_symlink() will honour the mapping gfp flags, so once you've done inode_nohighmem() it's safe to use, but if you allocate and insert the page manually, make sure to use the right gfp flags.

---

**mandatory**

->follow_link() is replaced with ->get_link(); same API, except that

- ->get_link() gets inode as a separate argument
- ->get_link() may be called in RCU mode - in that case NULL dentry is passed

---

**mandatory**

->get_link() gets struct delayed_call *`done` now, and should do set_delayed_call() where it used to set *`cookie`.

->put_link() is gone - just give the destructor to set_delayed_call() in ->get_link().

---

**mandatory**

->getxattr() and xattr_handler.get() get dentry and inode passed separately. dentry might be yet to be attached to inode, so do _not_ use its ->d_inode in the instances. Rationale: !@#!@# security_d_instantiate() needs to be called before we attach dentry to inode.

---

**mandatory**

symlinks are no longer the only inodes that do *not* have i_bdev/i_cdev/ i_pipe/i_link union zeroed out at inode eviction. As the result, you can't assume that non-NULL value in ->i_nlink at ->destroy_inode() implies that it's a symlink. Checking ->i_mode is really needed now. In-tree we had to fix shmem_destroy_callback() that used to take that kind of shortcut; watch out, since that shortcut is no longer valid.

---

**mandatory**

->i_mutex is replaced with ->i_rwsem now. inode_lock() et.al. work as they used to - they just take it exclusive. However, ->lookup() may be called with parent locked shared. Its instances must not

- use d_instantiate) and *d_rehash()* separately - use *d_add()* or *d_splice_alias()* instead.
- use *d_rehash()* alone - call d_add(new_dentry, NULL) instead.
- in the unlikely case when (read-only) access to filesystem data structures needs exclusion for some reason, arrange it yourself. None of the in-tree filesystems needed that.
- rely on ->d_parent and ->d_name not changing after dentry has been fed to *d_add()* or *d_splice_alias()*. Again, none of the in-tree instances relied upon that.

We are guaranteed that lookups of the same name in the same directory will not happen in parallel ("same" in the sense of your ->d_compare()). Lookups on different names in the same directory can and do happen in parallel now.

---

**mandatory**

->iterate_shared() is added. Exclusion on struct file level is still provided (as well as that between it and lseek on the same struct file), but if your directory has been opened several times, you can get these called in parallel. Exclusion between that method and all directory-modifying ones is still provided, of course.

If you have any per-inode or per-dentry in-core data structures modified by ->iterate_shared(), you might need something to serialize the access to them. If you do dcache pre-seeding, you'll need to switch to d_alloc_parallel() for that; look for in-tree examples.

---

**mandatory**

->atomic_open() calls without O_CREAT may happen in parallel.

---

**mandatory**

->setxattr() and xattr_handler.set() get dentry and inode passed separately. The xattr_handler.set() gets passed the user namespace of the mount the inode is seen from so

filesystems can idmap the i_uid and i_gid accordingly. dentry might be yet to be attached to inode, so do _not_ use its ->d_inode in the instances. Rationale: !@#!@# security_d_instantiate() needs to be called before we attach dentry to inode and !@#!@##!@$!$#!@#$!@$!@$ smack ->*d_instantiate()* uses not just ->getxattr() but ->setxattr() as well.

---

**mandatory**

->d_compare() doesn't get parent as a separate argument anymore. If you used it for finding the struct super_block involved, dentry->d_sb will work just as well; if it's something more complicated, use dentry->d_parent. Just be careful not to assume that fetching it more than once will yield the same value - in RCU mode it could change under you.

---

**mandatory**

->rename() has an added flags argument. Any flags not handled by the filesystem should result in EINVAL being returned.

---

**recommended**

->readlink is optional for symlinks. Don't set, unless filesystem needs to fake something for readlink(2).

---

**mandatory**

->getattr() is now passed a struct path rather than a vfsmount and dentry separately, and it now has request_mask and query_flags arguments to specify the fields and sync type requested by statx. Filesystems not supporting any statx-specific features may ignore the new arguments.

---

**mandatory**

->atomic_open() calling conventions have changed. Gone is `int *opened`, along with FILE_OPENED/FILE_CREATED. In place of those we have FMODE_OPENED/FMODE_CREATED, set in file->f_mode. Additionally, return value for 'called finish_no_open(), open it yourself' case has become 0, not 1. Since finish_no_open() itself is returning 0 now, that part does not need any changes in ->atomic_open() instances.

---

**mandatory**

alloc_file() has become static now; two wrappers are to be used instead. alloc_file_pseudo(inode, vfsmount, name, flags, ops) is for the cases when dentry needs to be created; that's the majority of old alloc_file() users. Calling conventions: on success a reference to new struct file is returned and callers reference to inode is subsumed by that. On failure, ERR_PTR() is returned and no caller's references are affected, so the caller needs to drop the inode reference it held. alloc_file_clone(file, flags, ops) does not affect any caller's references. On success you get a new struct file sharing the mount/dentry with the original, on failure - ERR_PTR().

---

**mandatory**

->clone_file_range() and ->dedupe_file_range have been replaced with ->remap_file_range(). See *Overview of the Linux Virtual File System* for more information.

---

**recommended**

->lookup() instances doing an equivalent of:

```
if (IS_ERR(inode))
        return ERR_CAST(inode);
return d_splice_alias(inode, dentry);
```

don't need to bother with the check - *d_splice_alias()* will do the right thing when given ERR_PTR(...) as inode. Moreover, passing NULL inode to *d_splice_alias()* will also do the right thing (equivalent of d_add(dentry, NULL); return NULL;), so that kind of special cases also doesn't need a separate treatment.

---

**strongly recommended**

take the RCU-delayed parts of ->destroy_inode() into a new method - ->free_inode(). If ->destroy_inode() becomes empty - all the better, just get rid of it. Synchronous work (e.g. the stuff that can't be done from an RCU callback, or any WARN_ON() where we want the stack trace) *might* be movable to ->evict_inode(); however, that goes only for the things that are not needed to balance something done by ->alloc_inode(). IOW, if it's cleaning up the stuff that might have accumulated over the life of in-core inode, ->evict_inode() might be a fit.

Rules for inode destruction:

- if ->destroy_inode() is non-NULL, it gets called
- if ->free_inode() is non-NULL, it gets scheduled by call_rcu()
- combination of NULL ->destroy_inode and NULL ->free_inode is treated as NULL/free_inode_nonrcu, to preserve the compatibility.

Note that the callback (be it via ->free_inode() or explicit call_rcu() in ->destroy_inode()) is *NOT* ordered wrt superblock destruction; as the matter of fact, the superblock and all associated structures might be already gone. The filesystem driver is guaranteed to be still there, but that's it. Freeing memory in the callback is fine; doing more than that is possible, but requires a lot of care and is best avoided.

---

**mandatory**

DCACHE_RCUACCESS is gone; having an RCU delay on dentry freeing is the default. DCACHE_NORCU opts out, and only d_alloc_pseudo() has any business doing so.

---

**mandatory**

d_alloc_pseudo() is internal-only; uses outside of alloc_file_pseudo() are very suspect (and won't work in modules). Such uses are very likely to be misspelled d_alloc_anon().

---

**mandatory**

[should've been added in 2016] stale comment in finish_open() nonwithstanding, failure exits in ->atomic_open() instances should *NOT* fput() the file, no matter what. Everything is handled by the caller.

---

**mandatory**

*clone_private_mount()* returns a longterm mount now, so the proper destructor of its result is kern_unmount() or kern_unmount_array().

---

**mandatory**

zero-length bvec segments are disallowed, they must be filtered out before passed on to an iterator.

---

**mandatory**

For bvec based itererators *bio_iov_iter_get_pages()* now doesn't copy bvecs but uses the one provided. Anyone issuing kiocb-I/O should ensure that the bvec and page references stay until I/O has completed, i.e. until ->ki_complete() has been called or returned with non -EIOCBQUEUED code.

---

**mandatory**

*mnt_want_write_file()* can now only be paired with mnt_drop_write_file(), whereas previously it could be paired with *mnt_drop_write()* as well.

---

**mandatory**

iov_iter_copy_from_user_atomic() is gone; use copy_page_from_iter_atomic(). The difference is copy_page_from_iter_atomic() advances the iterator and you don't need iov_iter_advance() after it. However, if you decide to use only a part of obtained data, you should do iov_iter_revert().

---

**mandatory**

Calling conventions for file_open_root() changed; now it takes struct path * instead of passing mount and dentry separately. For callers that used to pass <mnt, mnt->mnt_root> pair (i.e. the root of given mount), a new helper is provided - file_open_root_mnt(). In-tree users adjusted.

---

**mandatory**

no_llseek is gone; don't set .llseek to that - just leave it NULL instead. Checks for ”does that file have llseek(2), or should it fail with ESPIPE” should be done by looking at FMODE_LSEEK in file->f_mode.

---

*mandatory*

filldir_t (readdir callbacks) calling conventions have changed. Instead of returning 0 or -E... it returns bool now. false means "no more" (as -E... used to) and true - "keep going" (as 0 in old calling conventions). Rationale: callers never looked at specific -E... values anyway. -> iterate_shared() instances require no changes at all, all filldir_t ones in the tree converted.

---

**mandatory**

Calling conventions for ->tmpfile() have changed. It now takes a struct file pointer instead of struct dentry pointer. d_tmpfile() is similarly changed to simplify callers. The passed file is in a non-open state and on success must be opened before returning (e.g. by calling finish_open_simple()).

---

**mandatory**

Calling convention for ->huge_fault has changed. It now takes a page order instead of an enum page_entry_size, and it may be called without the mmap_lock held. All in-tree users have been audited and do not seem to depend on the mmap_lock being held, but out of tree users should verify for themselves. If they do need it, they can return VM_FAULT_RETRY to be called with the mmap_lock held.

---

**mandatory**

The order of opening block devices and matching or creating superblocks has changed.

The old logic opened block devices first and then tried to find a suitable superblock to reuse based on the block device pointer.

The new logic tries to find a suitable superblock first based on the device number, and opening the block device afterwards.

Since opening block devices cannot happen under s_umount because of lock ordering requirements s_umount is now dropped while opening block devices and reacquired before calling fill_super().

In the old logic concurrent mounters would find the superblock on the list of superblocks for the filesystem type. Since the first opener of the block device would hold s_umount they would wait until the superblock became either born or was discarded due to initialization failure.

Since the new logic drops s_umount concurrent mounters could grab s_umount and would spin. Instead they are now made to wait using an explicit wait-wake mechanism without having to hold s_umount.

---

**mandatory**

The holder of a block device is now the superblock.

The holder of a block device used to be the file_system_type which wasn't particularly useful. It wasn't possible to go from block device to owning superblock without matching on the device pointer stored in the superblock. This mechanism would only work for a single device so the block layer couldn't find the owning superblock of any additional devices.

In the old mechanism reusing or creating a superblock for a racing mount(2) and umount(2) relied on the file_system_type as the holder. This was severly underdocumented however:

(1) Any concurrent mounter that managed to grab an active reference on an existing superblock was made to wait until the superblock either became ready or until the superblock was removed from the list of superblocks of the filesystem type. If the superblock is ready the caller would simple reuse it.

(2) If the mounter came after *deactivate_locked_super()* but before the superblock had been removed from the list of superblocks of the filesystem type the mounter would wait until the superblock was shutdown, reuse the block device and allocate a new superblock.

(3) If the mounter came after *deactivate_locked_super()* and after the superblock had been removed from the list of superblocks of the filesystem type the mounter would reuse the block device and allocate a new superblock (the bd_holder point may still be set to the filesystem type).

Because the holder of the block device was the file_system_type any concurrent mounter could open the block devices of any superblock of the same file_system_type without risking seeing EBUSY because the block device was still in use by another superblock.

Making the superblock the owner of the block device changes this as the holder is now a unique superblock and thus block devices associated with it cannot be reused by concurrent mounters. So a concurrent mounter in (2) could suddenly see EBUSY when trying to open a block device whose holder was a different superblock.

The new logic thus waits until the superblock and the devices are shutdown in ->kill_sb(). Removal of the superblock from the list of superblocks of the filesystem type is now moved to a later point when the devices are closed:

(1) Any concurrent mounter managing to grab an active reference on an existing superblock is made to wait until the superblock is either ready or until the superblock and all devices are shutdown in ->kill_sb(). If the superblock is ready the caller will simply reuse it.

(2) If the mounter comes after *deactivate_locked_super()* but before the superblock has been removed from the list of superblocks of the filesystem type the mounter is made to wait until the superblock and the devices are shut down in ->kill_sb() and the superblock is removed from the list of superblocks of the filesystem type. The mounter will allocate a new superblock and grab ownership of the block device (the bd_holder pointer of the block device will be set to the newly allocated superblock).

(3) This case is now collapsed into (2) as the superblock is left on the list of superblocks of the filesystem type until all devices are shutdown in ->kill_sb(). In other words, if the superblock isn't on the list of superblock of the filesystem type anymore then it has given up ownership of all associated block devices (the bd_holder pointer is NULL).

As this is a VFS level change it has no practical consequences for filesystems other than that all of them must use one of the provided kill_litter_super(), kill_anon_super(), or kill_block_super() helpers.

---

**mandatory**

Lock ordering has been changed so that s_umount ranks above open_mutex again. All places where s_umount was taken under open_mutex have been fixed up.

---

**mandatory**

export_operations ->encode_fh() no longer has a default implementation to encode FILEID_INO32_GEN* file handles. Filesystems that used the default implementation may use the generic helper *generic_encode_ino32_fh()* explicitly.

---

**mandatory**

If ->rename() update of .. on cross-directory move needs an exclusion with directory modifications, do *not* lock the subdirectory in question in your ->rename() - it's done by the caller now [that item should've been added in 28eceeda130f "fs: Lock moved directories"].

---

**mandatory**

On same-directory ->rename() the (tautological) update of .. is not protected by any locks; just don't do it if the old parent is the same as the new one. We really can't lock two subdirectories in same-directory rename - not without deadlocks.

---

**mandatory**

lock_rename() and lock_rename_child() may fail in cross-directory case, if their arguments do not have a common ancestor. In that case ERR_PTR(-EXDEV) is returned, with no locks taken. In-tree users updated; out-of-tree ones would need to do so.

---

**mandatory**

The list of children anchored in parent dentry got turned into hlist now. Field names got changed (->d_children/->d_sib instead of ->d_subdirs/->d_child for anchor/entries resp.), so any affected places will be immediately caught by compiler.

---

**mandatory**

->*d_delete()* instances are now called for dentries with ->d_lock held and refcount equal to 0. They are not permitted to drop/regain ->d_lock. None of in-tree instances did anything of that sort. Make sure yours do not...

---

**mandatory**

->d_prune() instances are now called without ->d_lock held on the parent. ->d_lock on dentry itself is still held; if you need per-parent exclusions (none of the in-tree instances did), use your own spinlock.

->d_iput() and ->d_release() are called with victim dentry still in the list of parent's children. It is still unhashed, marked killed, etc., just not removed from parent's ->d_children yet.

Anyone iterating through the list of children needs to be aware of the half-killed dentries that might be seen there; taking ->d_lock on those will see them negative, unhashed and with negative refcount, which means that most of the in-kernel users would've done the right thing anyway without any adjustment.

---

**recommended**

Block device freezing and thawing have been moved to holder operations.

Before this change, get_active_super() would only be able to find the superblock of the main block device, i.e., the one stored in sb->s_bdev. Block device freezing now works for any block device owned by a given superblock, not just the main block device. The get_active_super() helper and bd_fsfreeze_sb pointer are gone.

CHAPTER

TWO

# FILESYSTEM SUPPORT LAYERS

Documentation for the support code within the filesystem layer for use in filesystem implementations.

## 2.1 The Linux Journalling API

### 2.1.1 Overview

#### Details

The journalling layer is easy to use. You need to first of all create a journal_t data structure. There are two calls to do this dependent on how you decide to allocate the physical media on which the journal resides. The *jbd2_journal_init_inode()* call is for journals stored in filesystem inodes, or the *jbd2_journal_init_dev()* call can be used for journal stored on a raw device (in a continuous range of blocks). A journal_t is a typedef for a struct pointer, so when you are finally finished make sure you call *jbd2_journal_destroy()* on it to free up any used kernel memory.

Once you have got your journal_t object you need to 'mount' or load the journal file. The journalling layer expects the space for the journal was already allocated and initialized properly by the userspace tools. When loading the journal you must call *jbd2_journal_load()* to process journal contents. If the client file system detects the journal contents does not need to be processed (or even need not have valid contents), it may call *jbd2_journal_wipe()* to clear the journal contents before calling *jbd2_journal_load()*.

Note that jbd2_journal_wipe(..,0) calls *jbd2_journal_skip_recovery()* for you if it detects any outstanding transactions in the journal and similarly *jbd2_journal_load()* will call *jbd2_journal_recover()* if necessary. I would advise reading ext4_load_journal() in fs/ext4/super.c for examples on this stage.

Now you can go ahead and start modifying the underlying filesystem. Almost.

You still need to actually journal your filesystem changes, this is done by wrapping them into transactions. Additionally you also need to wrap the modification of each of the buffers with calls to the journal layer, so it knows what the modifications you are actually making are. To do this use *jbd2_journal_start()* which returns a transaction handle.

*jbd2_journal_start()* and its counterpart *jbd2_journal_stop()*, which indicates the end of a transaction are nestable calls, so you can reenter a transaction if necessary, but remember you must call *jbd2_journal_stop()* the same number of times as *jbd2_journal_start()* before

the transaction is completed (or more accurately leaves the update phase). Ext4/VFS makes use of this feature to simplify handling of inode dirtying, quota support, etc.

Inside each transaction you need to wrap the modifications to the individual buffers (blocks). Before you start to modify a buffer you need to call *jbd2_journal_get_create_access()* / *jbd2_journal_get_write_access()* / *jbd2_journal_get_undo_access()* as appropriate, this allows the journalling layer to copy the unmodified data if it needs to. After all the buffer may be part of a previously uncommitted transaction. At this point you are at last ready to modify a buffer, and once you are have done so you need to call *jbd2_journal_dirty_metadata()*. Or if you've asked for access to a buffer you now know is now longer required to be pushed back on the device you can call *jbd2_journal_forget()* in much the same way as you might have used bforget() in the past.

A *jbd2_journal_flush()* may be called at any time to commit and checkpoint all your transactions.

Then at umount time , in your put_super() you can then call *jbd2_journal_destroy()* to clean up your in-core journal object.

Unfortunately there a couple of ways the journal layer can cause a deadlock. The first thing to note is that each task can only have a single outstanding transaction at any one time, remember nothing commits until the outermost *jbd2_journal_stop()*. This means you must complete the transaction at the end of each file/inode/address etc. operation you perform, so that the journalling system isn't re-entered on another journal. Since transactions can't be nested/batched across differing journals, and another filesystem other than yours (say ext4) may be modified in a later syscall.

The second case to bear in mind is that *jbd2_journal_start()* can block if there isn't enough space in the journal for your transaction (based on the passed nblocks param) - when it blocks it merely(!) needs to wait for transactions to complete and be committed from other tasks, so essentially we are waiting for *jbd2_journal_stop()*. So to avoid deadlocks you must treat *jbd2_journal_start()* / *jbd2_journal_stop()* as if they were semaphores and include them in your semaphore ordering rules to prevent deadlocks. Note that *jbd2_journal_extend()* has similar blocking behaviour to *jbd2_journal_start()* so you can deadlock here just as easily as on *jbd2_journal_start()*.

Try to reserve the right number of blocks the first time. ;-). This will be the maximum number of blocks you are going to touch in this transaction. I advise having a look at at least ext4_jbd.h to see the basis on which ext4 uses to make these decisions.

Another wriggle to watch out for is your on-disk block allocation strategy. Why? Because, if you do a delete, you need to ensure you haven't reused any of the freed blocks until the transaction freeing these blocks commits. If you reused these blocks and crash happens, there is no way to restore the contents of the reallocated blocks at the end of the last fully committed transaction. One simple way of doing this is to mark blocks as free in internal in-memory block allocation structures only after the transaction freeing them commits. Ext4 uses journal commit callback for this purpose.

With journal commit callbacks you can ask the journalling layer to call a callback function when the transaction is finally committed to disk, so that you can do some of your own management. You ask the journalling layer for calling the callback by simply setting `journal->j_commit_callback` function pointer and that function is called after each transaction commit. You can also use `transaction->t_private_list` for attaching entries to a transaction that need processing when the transaction commits.

JBD2 also provides a way to block all transaction updates via *jbd2_journal_lock_updates()*

/ *jbd2_journal_unlock_updates()*. Ext4 uses this when it wants a window with a clean and stable fs for a moment. E.g.

```
jbd2_journal_lock_updates() //stop new stuff happening..
jbd2_journal_flush()        // checkpoint everything.
..do stuff on stable fs
jbd2_journal_unlock_updates() // carry on with filesystem use.
```

The opportunities for abuse and DOS attacks with this should be obvious, if you allow unprivileged userspace to trigger codepaths containing these calls.

#### Fast commits

JBD2 to also allows you to perform file-system specific delta commits known as fast commits. In order to use fast commits, you will need to set following callbacks that perform correspodning work:

*journal->j_fc_cleanup_cb*: Cleanup function called after every full commit and fast commit.

*journal->j_fc_replay_cb*: Replay function called for replay of fast commit blocks.

File system is free to perform fast commits as and when it wants as long as it gets permission from JBD2 to do so by calling the function `jbd2_fc_begin_commit()`. Once a fast commit is done, the client file system should tell JBD2 about it by calling `jbd2_fc_end_commit()`. If file system wants JBD2 to perform a full commit immediately after stopping the fast commit it can do so by calling `jbd2_fc_end_commit_fallback()`. This is useful if fast commit operation fails for some reason and the only way to guarantee consistency is for JBD2 to perform the full traditional commit.

JBD2 helper functions to manage fast commit buffers. File system can use `jbd2_fc_get_buf()` and `jbd2_fc_wait_bufs()` to allocate and wait on IO completion of fast commit buffers.

Currently, only Ext4 implements fast commits. For details of its implementation of fast commits, please refer to the top level comments in fs/ext4/fast_commit.c.

#### Summary

Using the journal is a matter of wrapping the different context changes, being each mount, each modification (transaction) and each changed buffer to tell the journalling layer about them.

### 2.1.2 Data Types

The journalling layer uses typedefs to 'hide' the concrete definitions of the structures used. As a client of the JBD2 layer you can just rely on the using the pointer as a magic cookie of some sort. Obviously the hiding is not enforced as this is 'C'.

### Structures

type **handle_t**

The handle_t type represents a single atomic update being performed by some process.

**Description**

All filesystem modifications made by the process go through this handle. Recursive operations (such as quota operations) are gathered into a single update.

The buffer credits field is used to account for journaled buffers being modified by the running process. To ensure that there is enough log space for all outstanding operations, we need to limit the number of outstanding buffers possible at any time. When the operation completes, any buffer credits not used are credited back to the transaction, so that at all times we know how many buffers the outstanding updates on a transaction might possibly touch.

This is an opaque datatype.

type **journal_t**

The journal_t maintains all of the journaling state information for a single filesystem.

**Description**

journal_t is linked to from the fs superblock structure.

We use the journal_t to keep track of all outstanding transaction activity on the filesystem, and to manage the state of the log writing process.

This is an opaque datatype.

struct **jbd2_inode**

The jbd_inode type is the structure linking inodes in ordered mode present in a transaction so that we can sync them during commit.

**Definition**:

```
struct jbd2_inode {
    transaction_t *i_transaction;
    transaction_t *i_next_transaction;
    struct list_head i_list;
    struct inode *i_vfs_inode;
    unsigned long i_flags;
    loff_t i_dirty_start;
    loff_t i_dirty_end;
};
```

**Members**

**i_transaction**

Which transaction does this inode belong to? Either the running transaction or the committing one. [j_list_lock]

**i_next_transaction**

Pointer to the running transaction modifying inode's data in case there is already a committing transaction touching it. [j_list_lock]

**i_list**

List of inodes in the i_transaction [j_list_lock]

**i_vfs_inode**
VFS inode this inode belongs to [constant for lifetime of structure]

**i_flags**
Flags of inode [j_list_lock]

**i_dirty_start**
Offset in bytes where the dirty range for this inode starts. [j_list_lock]

**i_dirty_end**
Inclusive offset in bytes where the dirty range for this inode ends. [j_list_lock]

struct **jbd2_journal_handle**
The jbd2_journal_handle type is the concrete type associated with handle_t.

**Definition**:

```
struct jbd2_journal_handle {
    union {
        transaction_t *h_transaction;
        journal_t *h_journal;
    };
    handle_t *h_rsv_handle;
    int h_total_credits;
    int h_revoke_credits;
    int h_revoke_credits_requested;
    int h_ref;
    int h_err;
    unsigned int    h_sync:         1;
    unsigned int    h_jdata:        1;
    unsigned int    h_reserved:     1;
    unsigned int    h_aborted:      1;
    unsigned int    h_type:         8;
    unsigned int    h_line_no:      16;
    unsigned long           h_start_jiffies;
    unsigned int            h_requested_credits;
    unsigned int            saved_alloc_context;
};
```

**Members**

**{unnamed_union}**
anonymous

**h_transaction**
Which compound transaction is this update a part of?

**h_journal**
Which journal handle belongs to - used iff h_reserved set.

**h_rsv_handle**
Handle reserved for finishing the logical operation.

**h_total_credits**
Number of remaining buffers we are allowed to add to journal. These are dirty buffers and revoke descriptor blocks.

**h_revoke_credits**
Number of remaining revoke records available for handle

**h_revoke_credits_requested**
Holds **h_revoke_credits** after handle is started.

**h_ref**
Reference count on this handle.

**h_err**
Field for caller's use to track errors through large fs operations.

**h_sync**
Flag for sync-on-close.

**h_jdata**
Flag to force data journaling.

**h_reserved**
Flag for handle for reserved credits.

**h_aborted**
Flag indicating fatal error on handle.

**h_type**
For handle statistics.

**h_line_no**
For handle statistics.

**h_start_jiffies**
Handle Start time.

**h_requested_credits**
Holds **h_total_credits** after handle is started.

**saved_alloc_context**
Saved context while transaction is open.

struct **journal_s**
The journal_s type is the concrete type associated with journal_t.

**Definition**:

```
struct journal_s {
    unsigned long           j_flags;
    int j_errno;
    struct mutex            j_abort_mutex;
    struct buffer_head      *j_sb_buffer;
    journal_superblock_t *j_superblock;
    rwlock_t j_state_lock;
    int j_barrier_count;
    struct mutex            j_barrier;
    transaction_t *j_running_transaction;
    transaction_t *j_committing_transaction;
    transaction_t *j_checkpoint_transactions;
    wait_queue_head_t j_wait_transaction_locked;
    wait_queue_head_t j_wait_done_commit;
```

```
    wait_queue_head_t j_wait_commit;
    wait_queue_head_t j_wait_updates;
    wait_queue_head_t j_wait_reserved;
    wait_queue_head_t j_fc_wait;
    struct mutex            j_checkpoint_mutex;
    struct buffer_head      *j_chkpt_bhs[JBD2_NR_BATCH];
    struct shrinker         *j_shrinker;
    struct percpu_counter   j_checkpoint_jh_count;
    transaction_t *j_shrink_transaction;
    unsigned long           j_head;
    unsigned long           j_tail;
    unsigned long           j_free;
    unsigned long           j_first;
    unsigned long           j_last;
    unsigned long           j_fc_first;
    unsigned long           j_fc_off;
    unsigned long           j_fc_last;
    struct block_device     *j_dev;
    int j_blocksize;
    unsigned long long      j_blk_offset;
    char j_devname[BDEVNAME_SIZE+24];
    struct block_device     *j_fs_dev;
    errseq_t j_fs_dev_wb_err;
    unsigned int            j_total_len;
    atomic_t j_reserved_credits;
    spinlock_t j_list_lock;
    struct inode            *j_inode;
    tid_t j_tail_sequence;
    tid_t j_transaction_sequence;
    tid_t j_commit_sequence;
    tid_t j_commit_request;
    __u8 j_uuid[16];
    struct task_struct      *j_task;
    int j_max_transaction_buffers;
    int j_revoke_records_per_block;
    unsigned long           j_commit_interval;
    struct timer_list       j_commit_timer;
    spinlock_t j_revoke_lock;
    struct jbd2_revoke_table_s *j_revoke;
    struct jbd2_revoke_table_s *j_revoke_table[2];
    struct buffer_head      **j_wbuf;
    struct buffer_head      **j_fc_wbuf;
    int j_wbufsize;
    int j_fc_wbufsize;
    pid_t j_last_sync_writer;
    u64 j_average_commit_time;
    u32 j_min_batch_time;
    u32 j_max_batch_time;
    void (*j_commit_callback)(journal_t *, transaction_t *);
    int (*j_submit_inode_data_buffers) (struct jbd2_inode *);
```

```
    int (*j_finish_inode_data_buffers) (struct jbd2_inode *);
    spinlock_t j_history_lock;
    struct proc_dir_entry   *j_proc_entry;
    struct transaction_stats_s j_stats;
    unsigned int            j_failed_commit;
    void *j_private;
    struct crypto_shash *j_chksum_driver;
    __u32 j_csum_seed;
#ifdef CONFIG_DEBUG_LOCK_ALLOC;
    struct lockdep_map      j_trans_commit_map;
#endif;
    void (*j_fc_cleanup_callback)(struct journal_s *journal, int full, tid_t␣
↪tid);
    int (*j_fc_replay_callback)(struct journal_s *journal,struct buffer_head␣
↪*bh,enum passtype pass, int off, tid_t expected_commit_id);
    int (*j_bmap)(struct journal_s *journal, sector_t *block);
};
```

**Members**

**j_flags**
General journaling state flags [j_state_lock, no lock for quick racy checks]

**j_errno**
Is there an outstanding uncleared error on the journal (from a prior abort)? [j_state_lock]

**j_abort_mutex**
Lock the whole aborting procedure.

**j_sb_buffer**
The first part of the superblock buffer.

**j_superblock**
The second part of the superblock buffer.

**j_state_lock**
Protect the various scalars in the journal.

**j_barrier_count**
Number of processes waiting to create a barrier lock [j_state_lock, no lock for quick racy checks]

**j_barrier**
The barrier lock itself.

**j_running_transaction**
Transactions: The current running transaction... [j_state_lock, no lock for quick racy checks] [caller holding open handle]

**j_committing_transaction**
the transaction we are pushing to disk [j_state_lock] [caller holding open handle]

**j_checkpoint_transactions**
... and a linked circular list of all transactions waiting for checkpointing. [j_list_lock]

**j_wait_transaction_locked**
Wait queue for waiting for a locked transaction to start committing, or for a barrier lock

to be released.

**j_wait_done_commit**
Wait queue for waiting for commit to complete.

**j_wait_commit**
Wait queue to trigger commit.

**j_wait_updates**
Wait queue to wait for updates to complete.

**j_wait_reserved**
Wait queue to wait for reserved buffer credits to drop.

**j_fc_wait**
Wait queue to wait for completion of async fast commits.

**j_checkpoint_mutex**
Semaphore for locking against concurrent checkpoints.

**j_chkpt_bhs**
List of buffer heads used by the checkpoint routine. This was moved from jbd2_log_do_checkpoint() to reduce stack usage. Access to this array is controlled by the **j_checkpoint_mutex**. [j_checkpoint_mutex]

**j_shrinker**
Journal head shrinker, reclaim buffer's journal head which has been written back.

**j_checkpoint_jh_count**
Number of journal buffers on the checkpoint list. [j_list_lock]

**j_shrink_transaction**
Record next transaction will shrink on the checkpoint list. [j_list_lock]

**j_head**
Journal head: identifies the first unused block in the journal. [j_state_lock]

**j_tail**
Journal tail: identifies the oldest still-used block in the journal. [j_state_lock]

**j_free**
Journal free: how many free blocks are there in the journal? [j_state_lock]

**j_first**
The block number of the first usable block in the journal [j_state_lock].

**j_last**
The block number one beyond the last usable block in the journal [j_state_lock].

**j_fc_first**
The block number of the first fast commit block in the journal [j_state_lock].

**j_fc_off**
Number of fast commit blocks currently allocated. Accessed only during fast commit. Currently only process can do fast commit, so this field is not protected by any lock.

**j_fc_last**
The block number one beyond the last fast commit block in the journal [j_state_lock].

**j_dev**
Device where we store the journal.

**j_blocksize**
Block size for the location where we store the journal.

**j_blk_offset**
Starting block offset into the device where we store the journal.

**j_devname**
Journal device name.

**j_fs_dev**
Device which holds the client fs. For internal journal this will be equal to j_dev.

**j_fs_dev_wb_err**
Records the errseq of the client fs's backing block device.

**j_total_len**
Total maximum capacity of the journal region on disk.

**j_reserved_credits**
Number of buffers reserved from the running transaction.

**j_list_lock**
Protects the buffer lists and internal buffer state.

**j_inode**
Optional inode where we store the journal. If present, all journal block numbers are mapped into this inode via *bmap()*.

**j_tail_sequence**
Sequence number of the oldest transaction in the log [j_state_lock]

**j_transaction_sequence**
Sequence number of the next transaction to grant [j_state_lock]

**j_commit_sequence**
Sequence number of the most recently committed transaction [j_state_lock, no lock for quick racy checks]

**j_commit_request**
Sequence number of the most recent transaction wanting commit [j_state_lock, no lock for quick racy checks]

**j_uuid**
Journal uuid: identifies the object (filesystem, LVM volume etc) backed by this journal. This will eventually be replaced by an array of uuids, allowing us to index multiple devices within a single journal and to perform atomic updates across them.

**j_task**
Pointer to the current commit thread for this journal.

**j_max_transaction_buffers**
Maximum number of metadata buffers to allow in a single compound commit transaction.

**j_revoke_records_per_block**
Number of revoke records that fit in one descriptor block.

**j_commit_interval**
What is the maximum transaction lifetime before we begin a commit?

**j_commit_timer**
The timer used to wakeup the commit thread.

**j_revoke_lock**
Protect the revoke table.

**j_revoke**
The revoke table - maintains the list of revoked blocks in the current transaction.

**j_revoke_table**
Alternate revoke tables for j_revoke.

**j_wbuf**
Array of bhs for jbd2_journal_commit_transaction.

**j_fc_wbuf**
Array of fast commit bhs for fast commit. Accessed only during a fast commit. Currently only process can do fast commit, so this field is not protected by any lock.

**j_wbufsize**
Size of **j_wbuf** array.

**j_fc_wbufsize**
Size of **j_fc_wbuf** array.

**j_last_sync_writer**
The pid of the last person to run a synchronous operation through the journal.

**j_average_commit_time**
The average amount of time in nanoseconds it takes to commit a transaction to disk. [j_state_lock]

**j_min_batch_time**
Minimum time that we should wait for additional filesystem operations to get batched into a synchronous handle in microseconds.

**j_max_batch_time**
Maximum time that we should wait for additional filesystem operations to get batched into a synchronous handle in microseconds.

**j_commit_callback**
This function is called when a transaction is closed.

**j_submit_inode_data_buffers**
This function is called for all inodes associated with the committing transaction marked with JI_WRITE_DATA flag before we start to write out the transaction to the journal.

**j_finish_inode_data_buffers**
This function is called for all inodes associated with the committing transaction marked with JI_WAIT_DATA flag after we have written the transaction to the journal but before we write out the commit block.

**j_history_lock**
Protect the transactions statistics history.

**j_proc_entry**
procfs entry for the jbd statistics directory.

**j_stats**
Overall statistics.

**j_failed_commit**
Failed journal commit ID.

**j_private**
An opaque pointer to fs-private information. ext3 puts its superblock pointer here.

**j_chksum_driver**
Reference to checksum algorithm driver via cryptoapi.

**j_csum_seed**
Precomputed journal UUID checksum for seeding other checksums.

**j_trans_commit_map**
Lockdep entity to track transaction commit dependencies. Handles hold this "lock" for read, when we wait for commit, we acquire the "lock" for writing. This matches the properties of jbd2 journalling where the running transaction has to wait for all handles to be dropped to commit that transaction and also acquiring a handle may require transaction commit to finish.

**j_fc_cleanup_callback**
Clean-up after fast commit or full commit. JBD2 calls this function after every commit operation.

**j_fc_replay_callback**
File-system specific function that performs replay of a fast commit. JBD2 calls this function for each fast commit block found in the journal. This function should return JBD2_FC_REPLAY_CONTINUE to indicate that the block was processed correctly and more fast commit replay should continue. Return value of JBD2_FC_REPLAY_STOP indicates the end of replay (no more blocks remaining). A negative return value indicates error.

**j_bmap**
Bmap function that should be used instead of the generic VFS bmap function.

### 2.1.3 Functions

The functions here are split into two groups those that affect a journal as a whole, and those which are used to manage transactions

#### Journal Level

int **jbd2_journal_force_commit_nested**(*journal_t* *journal)
Force and wait upon a commit if the calling process is not within transaction.

**Parameters**

**journal_t *journal**
journal to force Returns true if progress was made.

**Description**

This is used for forcing out undo-protected data which contains bitmaps, when the fs is running out of space.

int **jbd2_journal_force_commit**(*journal_t* *journal)
force any uncommitted transactions

**Parameters**

**journal_t *journal**
journal to force

**Description**

Caller want unconditional commit. We can only force the running transaction if we don't have an active handle, otherwise, we will deadlock.

*journal_t* ***jbd2_journal_init_dev**(struct block_device *bdev, struct block_device *fs_dev, unsigned long long start, int len, int blocksize)

creates and initialises a journal structure

**Parameters**

**struct block_device *bdev**
Block device on which to create the journal

**struct block_device *fs_dev**
Device which hold journalled filesystem for this journal.

**unsigned long long start**
Block nr Start of journal.

**int len**
Length of the journal in blocks.

**int blocksize**
blocksize of journalling device

**Return**

a newly created journal_t *

jbd2_journal_init_dev creates a journal which maps a fixed contiguous range of blocks on an arbitrary block device.

*journal_t* ***jbd2_journal_init_inode**(struct *inode* *inode)

creates a journal which maps to a inode.

**Parameters**

**struct inode *inode**
An inode to create the journal in

**Description**

jbd2_journal_init_inode creates a journal which maps an on-disk inode as the journal. The inode must exist already, must support *bmap()* and must have all data blocks preallocated.

void **jbd2_journal_update_sb_errno**(*journal_t* *journal)

Update error in the journal.

**Parameters**

**journal_t *journal**
The journal to update.

**Description**

Update a journal's errno. Write updated superblock to disk waiting for IO to complete.

int **jbd2_journal_load**(*journal_t* *journal)

Read journal from disk.

**Parameters**

**journal_t *journal**
Journal to act on.

**Description**

Given a journal_t structure which tells us which disk blocks contain a journal, read the journal from disk to initialise the in-memory structures.

int **jbd2_journal_destroy**(*journal_t* *journal)

Release a journal_t structure.

**Parameters**

**journal_t *journal**
Journal to act on.

**Description**

Release a journal_t structure once it is no longer in use by the journaled object. Return <0 if we couldn't clean up the journal.

int **jbd2_journal_check_used_features**(*journal_t* *journal, unsigned long compat, unsigned long ro, unsigned long incompat)

Check if features specified are used.

**Parameters**

**journal_t *journal**
Journal to check.

**unsigned long compat**
bitmask of compatible features

**unsigned long ro**
bitmask of features that force read-only mount

**unsigned long incompat**
bitmask of incompatible features

**Description**

Check whether the journal uses all of a given set of features. Return true (non-zero) if it does.

int **jbd2_journal_check_available_features**(*journal_t* *journal, unsigned long compat, unsigned long ro, unsigned long incompat)

Check feature set in journalling layer

**Parameters**

**journal_t *journal**
Journal to check.

**unsigned long compat**
bitmask of compatible features

**unsigned long ro**
bitmask of features that force read-only mount

**unsigned long incompat**
bitmask of incompatible features

**Description**

Check whether the journaling code supports the use of all of a given set of features on this journal. Return true

int **jbd2_journal_set_features**(*journal_t* *journal, unsigned long compat, unsigned long ro, unsigned long incompat)

Mark a given journal feature in the superblock

**Parameters**

**journal_t *journal**
Journal to act on.

**unsigned long compat**
bitmask of compatible features

**unsigned long ro**
bitmask of features that force read-only mount

**unsigned long incompat**
bitmask of incompatible features

**Description**

Mark a given journal feature as present on the superblock. Returns true if the requested features could be set.

int **jbd2_journal_flush**(*journal_t* *journal, unsigned int flags)

Flush journal

**Parameters**

**journal_t *journal**
Journal to act on.

**unsigned int flags**
optional operation on the journal blocks after the flush (see below)

**Description**

Flush all data for a given journal to disk and empty the journal. Filesystems can use this when remounting readonly to ensure that recovery does not need to happen on remount. Optionally, a discard or zeroout can be issued on the journal blocks after flushing.

**flags:**
JBD2_JOURNAL_FLUSH_DISCARD: issues discards for the journal blocks JBD2_JOURNAL_FLUSH_ZEROOUT: issues zeroouts for the journal blocks

int **jbd2_journal_wipe**(*journal_t* *journal, int write)

Wipe journal contents

**Parameters**

**journal_t *journal**
Journal to act on.

**int write**
flag (see below)

**Description**

Wipe out all of the contents of a journal, safely. This will produce a warning if the journal contains any valid recovery information. Must be called between journal_init_*() and *jbd2_journal_load()*.

If 'write' is non-zero, then we wipe out the journal on disk; otherwise we merely suppress recovery.

void **jbd2_journal_abort**(*journal_t* *journal, int errno)
Shutdown the journal immediately.

**Parameters**

**journal_t *journal**
the journal to shutdown.

**int errno**
an error number to record in the journal indicating the reason for the shutdown.

**Description**

Perform a complete, immediate shutdown of the ENTIRE journal (not of a single transaction). This operation cannot be undone without closing and reopening the journal.

The jbd2_journal_abort function is intended to support higher level error recovery mechanisms such as the ext2/ext3 remount-readonly error mode.

Journal abort has very specific semantics. Any existing dirty, unjournaled buffers in the main filesystem will still be written to disk by bdflush, but the journaling mechanism will be suspended immediately and no further transaction commits will be honoured.

Any dirty, journaled buffers will be written back to disk without hitting the journal. Atomicity cannot be guaranteed on an aborted filesystem, but we _do_ attempt to leave as much data as possible behind for fsck to use for cleanup.

Any attempt to get a new transaction handle on a journal which is in ABORT state will just result in an -EROFS error return. A jbd2_journal_stop on an existing handle will return -EIO if we have entered abort state during the update.

Recursive transactions are not disturbed by journal abort until the final jbd2_journal_stop, which will receive the -EIO error.

Finally, the jbd2_journal_abort call allows the caller to supply an errno which will be recorded (if possible) in the journal superblock. This allows a client to record failure conditions in the middle of a transaction without having to complete the transaction to record the failure to disk. ext3_error, for example, now uses this functionality.

int **jbd2_journal_errno**(*journal_t* *journal)
returns the journal's error state.

**Parameters**

**journal_t *journal**
journal to examine.

**Description**

This is the errno number set with *jbd2_journal_abort()*, the last time the journal was mounted - if the journal was stopped without calling abort this will be 0.

If the journal has been aborted on this mount time -EROFS will be returned.

int **jbd2_journal_clear_err**(*journal_t* *journal)
clears the journal's error state

**Parameters**

**journal_t *journal**
journal to act on.

**Description**

An error must be cleared or acked to take a FS out of readonly mode.

void **jbd2_journal_ack_err**(*journal_t* *journal)
Ack journal err.

**Parameters**

**journal_t *journal**
journal to act on.

**Description**

An error must be cleared or acked to take a FS out of readonly mode.

int **jbd2_journal_recover**(*journal_t* *journal)
recovers a on-disk journal

**Parameters**

**journal_t *journal**
the journal to recover

**Description**

The primary function for recovering the log contents when mounting a journaled device.

Recovery is done in three passes. In the first pass, we look for the end of the log. In the second, we assemble the list of revoke blocks. In the third and final pass, we replay any un-revoked blocks in the log.

int **jbd2_journal_skip_recovery**(*journal_t* *journal)
Start journal and wipe exiting records

**Parameters**

**journal_t *journal**
journal to startup

**Description**

Locate any valid recovery information from the journal and set up the journal structures in memory to ignore it (presumably because the caller has evidence that it is out of date). This function doesn't appear to be exported..

We perform one pass over the journal to allow us to tell the user how much recovery information is being erased, and to let us initialise the journal transaction sequence numbers to the next unused ID.

### Transasction Level

*handle_t* ***jbd2_journal_start**(*journal_t* *journal, int nblocks)

Obtain a new handle.

**Parameters**

**journal_t *journal**
Journal to start transaction on.

**int nblocks**
number of block buffer we might modify

**Description**

We make sure that the transaction can guarantee at least nblocks of modified buffers in the log. We block until the log can guarantee that much space. Additionally, if rsv_blocks > 0, we also create another handle with rsv_blocks reserved blocks in the journal. This handle is stored in h_rsv_handle. It is not attached to any particular transaction and thus doesn't block transaction commit. If the caller uses this reserved handle, it has to set h_rsv_handle to NULL as otherwise *jbd2_journal_stop()* on the parent handle will dispose the reserved one. Reserved handle has to be converted to a normal handle using *jbd2_journal_start_reserved()* before it can be used.

Return a pointer to a newly allocated handle, or an ERR_PTR() value on failure.

int **jbd2_journal_start_reserved**(*handle_t* *handle, unsigned int type, unsigned int line_no)

start reserved handle

**Parameters**

**handle_t *handle**
handle to start

**unsigned int type**
for handle statistics

**unsigned int line_no**
for handle statistics

**Description**

Start handle that has been previously reserved with jbd2_journal_reserve(). This attaches **handle** to the running transaction (or creates one if there's not transaction running). Unlike *jbd2_journal_start()* this function cannot block on journal commit, checkpointing, or similar stuff. It can block on memory allocation or frozen journal though.

Return 0 on success, non-zero on error - handle is freed in that case.

int **jbd2_journal_extend**(*handle_t* *handle, int nblocks, int revoke_records)

extend buffer credits.

**Parameters**

**handle_t *handle**
handle to 'extend'

**int nblocks**
nr blocks to try to extend by.

**int revoke_records**
number of revoke records to try to extend by.

**Description**

Some transactions, such as large extends and truncates, can be done atomically all at once or in several stages. The operation requests a credit for a number of buffer modifications in advance, but can extend its credit if it needs more.

jbd2_journal_extend tries to give the running handle more buffer credits. It does not guarantee that allocation - this is a best-effort only. The calling process MUST be able to deal cleanly with a failure to extend here.

Return 0 on success, non-zero on failure.

return code < 0 implies an error return code > 0 implies normal transaction-full status.

int **jbd2__journal_restart**(*handle_t* *handle, int nblocks, int revoke_records, gfp_t gfp_mask)

restart a handle .

**Parameters**

**handle_t *handle**
handle to restart

**int nblocks**
nr credits requested

**int revoke_records**
number of revoke record credits requested

**gfp_t gfp_mask**
memory allocation flags (for start_this_handle)

**Description**

Restart a handle for a multi-transaction filesystem operation.

If the *jbd2_journal_extend()* call above fails to grant new buffer credits to a running handle, a call to jbd2_journal_restart will commit the handle's transaction so far and reattach the handle to a new transaction capable of guaranteeing the requested number of credits. We preserve reserved handle if there's any attached to the passed in handle.

void **jbd2_journal_lock_updates**(*journal_t* *journal)

establish a transaction barrier.

**Parameters**

**journal_t *journal**
Journal to establish a barrier on.

**Description**

This locks out any further updates from being started, and blocks until all existing updates have completed, returning only once the journal is in a quiescent state with no updates running.

The journal lock should not be held on entry.

void **jbd2_journal_unlock_updates**(*journal_t* *journal)
release barrier

**Parameters**

**journal_t *journal**
Journal to release the barrier on.

**Description**

Release a transaction barrier obtained with *jbd2_journal_lock_updates()*.

Should be called without the journal lock held.

int **jbd2_journal_get_write_access**(*handle_t* *handle, struct buffer_head *bh)
notify intent to modify a buffer for metadata (not data) update.

**Parameters**

**handle_t *handle**
transaction to add buffer modifications to

**struct buffer_head *bh**
bh to be used for metadata writes

**Return**

error code or 0 on success.

**Description**

In full data journalling mode the buffer may be of type BJ_AsyncData, because we're `write()ing` a buffer which is also part of a shared mapping.

int **jbd2_journal_get_create_access**(*handle_t* *handle, struct buffer_head *bh)
notify intent to use newly created bh

**Parameters**

**handle_t *handle**
transaction to new buffer to

**struct buffer_head *bh**
new buffer.

**Description**

Call this if you create a new bh.

int **jbd2_journal_get_undo_access**(*handle_t* *handle, struct buffer_head *bh)
Notify intent to modify metadata with non-rewindable consequences

**Parameters**

**handle_t *handle**
transaction

**struct buffer_head *bh**
buffer to undo

**Description**

Sometimes there is a need to distinguish between metadata which has been committed to disk and that which has not. The ext3fs code uses this for freeing and allocating space, we have to make sure that we do not reuse freed space until the deallocation has been committed, since if we overwrote that space we would make the delete un-rewindable in case of a crash.

To deal with that, jbd2_journal_get_undo_access requests write access to a buffer for parts of non-rewindable operations such as delete operations on the bitmaps. The journaling code must keep a copy of the buffer's contents prior to the undo_access call until such time as we know that the buffer has definitely been committed to disk.

We never need to know which transaction the committed data is part of, buffers touched here are guaranteed to be dirtied later and so will be committed to a new transaction in due course, at which point we can discard the old committed data pointer.

Returns error number or 0 on success.

void **jbd2_journal_set_triggers**(struct buffer_head *bh, struct jbd2_buffer_trigger_type *type)
Add triggers for commit writeout

**Parameters**

**struct buffer_head *bh**
buffer to trigger on

**struct jbd2_buffer_trigger_type *type**
struct jbd2_buffer_trigger_type containing the trigger(s).

**Description**

Set any triggers on this journal_head. This is always safe, because triggers for a committing buffer will be saved off, and triggers for a running transaction will match the buffer in that transaction.

Call with NULL to clear the triggers.

int **jbd2_journal_dirty_metadata**(*handle_t* *handle, struct buffer_head *bh)
mark a buffer as containing dirty metadata

**Parameters**

**handle_t *handle**
transaction to add buffer to.

**struct buffer_head *bh**
buffer to mark

**Description**

mark dirty metadata which needs to be journaled as part of the current transaction.

The buffer must have previously had *jbd2_journal_get_write_access()* called so that it has a valid journal_head attached to the buffer head.

The buffer is placed on the transaction's metadata list and is marked as belonging to the transaction.

Returns error number or 0 on success.

Special care needs to be taken if the buffer already belongs to the current committing transaction (in which case we should have frozen data present for that commit). In that case, we don't relink the buffer: that only gets done when the old transaction finally completes its commit.

int **jbd2_journal_forget**(*handle_t* *handle, struct buffer_head *bh)

bforget() for potentially-journaled buffers.

**Parameters**

**handle_t *handle**
transaction handle

**struct buffer_head *bh**
bh to 'forget'

**Description**

We can only do the bforget if there are no commits pending against the buffer. If the buffer is dirty in the current running transaction we can safely unlink it.

bh may not be a journalled buffer at all - it may be a non-JBD buffer which came off the hashtable. Check for this.

Decrements bh->b_count by one.

Allow this call even if the handle has aborted --- it may be part of the caller's cleanup after an abort.

int **jbd2_journal_stop**(*handle_t* *handle)

complete a transaction

**Parameters**

**handle_t *handle**
transaction to complete.

**Description**

All done for a particular handle.

There is not much action needed here. We just return any remaining buffer credits to the transaction and remove the handle. The only complication is that we need to start a commit operation if the filesystem is marked for synchronous update.

jbd2_journal_stop itself will not usually return an error, but it may do so in unusual circumstances. In particular, expect it to return -EIO if a jbd2_journal_abort has been executed since the transaction began.

bool **jbd2_journal_try_to_free_buffers**(*journal_t* *journal, struct *folio* *folio)

try to free page buffers.

**Parameters**

**journal_t *journal**
journal for operation

**struct folio *folio**
Folio to detach data from.

### Description

For all the buffers on this page, if they are fully written out ordered data, move them onto BUF_CLEAN so try_to_free_buffers() can reap them.

This function returns non-zero if we wish try_to_free_buffers() to be called. We do this if the page is releasable by try_to_free_buffers(). We also do it if the page has locked or dirty buffers and the caller wants us to perform sync or async writeout.

This complicates JBD locking somewhat. We aren't protected by the BKL here. We wish to remove the buffer from its committing or running transaction's ->t_datalist via __jbd2_journal_unfile_buffer.

This may *change* the value of transaction_t->t_datalist, so anyone who looks at t_datalist needs to lock against this function.

Even worse, someone may be doing a jbd2_journal_dirty_data on this buffer. So we need to lock against that. jbd2_journal_dirty_data() will come out of the lock with the buffer dirty, which makes it ineligible for release here.

Who else is affected by this? hmm... Really the only contender is do_get_write_access() - it could be looking at the buffer while journal_try_to_free_buffer() is changing its state. But that cannot happen because we never reallocate freed data as metadata while the data is part of a transaction. Yes?

Return false on failure, true on success

int **jbd2_journal_invalidate_folio**(*journal_t* *journal, struct *folio* *folio, size_t offset, size_t length)

### Parameters

**journal_t *journal**
: journal to use for flush...

**struct folio *folio**
: folio to flush

**size_t offset**
: start of the range to invalidate

**size_t length**
: length of the range to invalidate

### Description

Reap page buffers containing data after in the specified range in page. Can return -EBUSY if buffers are part of the committing transaction and the page is straddling i_size. Caller then has to wait for current commit and try again.

### 2.1.4 See also

Journaling the Linux ext2fs Filesystem, LinuxExpo 98, Stephen Tweedie

Ext3 Journalling FileSystem, OLS 2000, Dr. Stephen Tweedie

## 2.2 Filesystem-level encryption (fscrypt)

### 2.2.1 Introduction

fscrypt is a library which filesystems can hook into to support transparent encryption of files and directories.

Note: "fscrypt" in this document refers to the kernel-level portion, implemented in `fs/crypto/`, as opposed to the userspace tool fscrypt. This document only covers the kernel-level portion. For command-line examples of how to use encryption, see the documentation for the userspace tool fscrypt. Also, it is recommended to use the fscrypt userspace tool, or other existing userspace tools such as fscryptctl or Android's key management system, over using the kernel's API directly. Using existing tools reduces the chance of introducing your own security bugs. (Nevertheless, for completeness this documentation covers the kernel's API anyway.)

Unlike dm-crypt, fscrypt operates at the filesystem level rather than at the block device level. This allows it to encrypt different files with different keys and to have unencrypted files on the same filesystem. This is useful for multi-user systems where each user's data-at-rest needs to be cryptographically isolated from the others. However, except for filenames, fscrypt does not encrypt filesystem metadata.

Unlike eCryptfs, which is a stacked filesystem, fscrypt is integrated directly into supported filesystems --- currently ext4, F2FS, UBIFS, and CephFS. This allows encrypted files to be read and written without caching both the decrypted and encrypted pages in the pagecache, thereby nearly halving the memory used and bringing it in line with unencrypted files. Similarly, half as many dentries and inodes are needed. eCryptfs also limits encrypted filenames to 143 bytes, causing application compatibility issues; fscrypt allows the full 255 bytes (NAME_MAX). Finally, unlike eCryptfs, the fscrypt API can be used by unprivileged users, with no need to mount anything.

fscrypt does not support encrypting files in-place. Instead, it supports marking an empty directory as encrypted. Then, after userspace provides the key, all regular files, directories, and symbolic links created in that directory tree are transparently encrypted.

### 2.2.2 Threat model

#### Offline attacks

Provided that userspace chooses a strong encryption key, fscrypt protects the confidentiality of file contents and filenames in the event of a single point-in-time permanent offline compromise of the block device content. fscrypt does not protect the confidentiality of non-filename metadata, e.g. file sizes, file permissions, file timestamps, and extended attributes. Also, the existence and location of holes (unallocated blocks which logically contain all zeroes) in files is not protected.

fscrypt is not guaranteed to protect confidentiality or authenticity if an attacker is able to manipulate the filesystem offline prior to an authorized user later accessing the filesystem.

#### Online attacks

fscrypt (and storage encryption in general) can only provide limited protection, if any at all, against online attacks. In detail:

#### Side-channel attacks

fscrypt is only resistant to side-channel attacks, such as timing or electromagnetic attacks, to the extent that the underlying Linux Cryptographic API algorithms or inline encryption hardware are. If a vulnerable algorithm is used, such as a table-based implementation of AES, it may be possible for an attacker to mount a side channel attack against the online system. Side channel attacks may also be mounted against applications consuming decrypted data.

#### Unauthorized file access

After an encryption key has been added, fscrypt does not hide the plaintext file contents or filenames from other users on the same system. Instead, existing access control mechanisms such as file mode bits, POSIX ACLs, LSMs, or namespaces should be used for this purpose.

(For the reasoning behind this, understand that while the key is added, the confidentiality of the data, from the perspective of the system itself, is *not* protected by the mathematical properties of encryption but rather only by the correctness of the kernel. Therefore, any encryption-specific access control checks would merely be enforced by kernel *code* and therefore would be largely redundant with the wide variety of access control mechanisms already available.)

#### Kernel memory compromise

An attacker who compromises the system enough to read from arbitrary memory, e.g. by mounting a physical attack or by exploiting a kernel security vulnerability, can compromise all encryption keys that are currently in use.

However, fscrypt allows encryption keys to be removed from the kernel, which may protect them from later compromise.

In more detail, the FS_IOC_REMOVE_ENCRYPTION_KEY ioctl (or the FS_IOC_REMOVE_ENCRYPTION_KEY_ALL_USERS ioctl) can wipe a master encryption key from kernel memory. If it does so, it will also try to evict all cached inodes which had been ”unlocked” using the key, thereby wiping their per-file keys and making them once again appear ”locked”, i.e. in ciphertext or encrypted form.

However, these ioctls have some limitations:

- Per-file keys for in-use files will *not* be removed or wiped. Therefore, for maximum effect, userspace should close the relevant encrypted files and directories before removing a master key, as well as kill any processes whose working directory is in an affected encrypted directory.

- The kernel cannot magically wipe copies of the master key(s) that userspace might have as well. Therefore, userspace must wipe all copies of the master key(s) it makes as well; normally this should be done immediately after FS_IOC_ADD_ENCRYPTION_KEY, without waiting for FS_IOC_REMOVE_ENCRYPTION_KEY. Naturally, the same also applies to all higher levels in the key hierarchy. Userspace should also follow other security precautions such as mlock()ing memory containing keys to prevent it from being swapped out.
- In general, decrypted contents and filenames in the kernel VFS caches are freed but not wiped. Therefore, portions thereof may be recoverable from freed memory, even after the corresponding key(s) were wiped. To partially solve this, you can set CONFIG_PAGE_POISONING=y in your kernel config and add page_poison=1 to your kernel command line. However, this has a performance cost.
- Secret keys might still exist in CPU registers, in crypto accelerator hardware (if used by the crypto API to implement any of the algorithms), or in other places not explicitly considered here.

#### Limitations of v1 policies

v1 encryption policies have some weaknesses with respect to online attacks:

- There is no verification that the provided master key is correct. Therefore, a malicious user can temporarily associate the wrong key with another user's encrypted files to which they have read-only access. Because of filesystem caching, the wrong key will then be used by the other user's accesses to those files, even if the other user has the correct key in their own keyring. This violates the meaning of ”read-only access”.
- A compromise of a per-file key also compromises the master key from which it was derived.
- Non-root users cannot securely remove encryption keys.

All the above problems are fixed with v2 encryption policies. For this reason among others, it is recommended to use v2 encryption policies on all new encrypted directories.

### 2.2.3 Key hierarchy

#### Master Keys

Each encrypted directory tree is protected by a *master key*. Master keys can be up to 64 bytes long, and must be at least as long as the greater of the security strength of the contents and filenames encryption modes being used. For example, if any AES-256 mode is used, the master key must be at least 256 bits, i.e. 32 bytes. A stricter requirement applies if the key is used by a v1 encryption policy and AES-256-XTS is used; such keys must be 64 bytes.

To ”unlock” an encrypted directory tree, userspace must provide the appropriate master key. There can be any number of master keys, each of which protects any number of directory trees on any number of filesystems.

Master keys must be real cryptographic keys, i.e. indistinguishable from random bytestrings of the same length. This implies that users **must not** directly use a password as a master key, zero-pad a shorter key, or repeat a shorter key. Security cannot be guaranteed if userspace makes any such error, as the cryptographic proofs and analysis would no longer apply.

Instead, users should generate master keys either using a cryptographically secure random number generator, or by using a KDF (Key Derivation Function). The kernel does not do any key stretching; therefore, if userspace derives the key from a low-entropy secret such as a passphrase, it is critical that a KDF designed for this purpose be used, such as scrypt, PBKDF2, or Argon2.

### Key derivation function

With one exception, fscrypt never uses the master key(s) for encryption directly. Instead, they are only used as input to a KDF (Key Derivation Function) to derive the actual keys.

The KDF used for a particular master key differs depending on whether the key is used for v1 encryption policies or for v2 encryption policies. Users **must not** use the same key for both v1 and v2 encryption policies. (No real-world attack is currently known on this specific case of key reuse, but its security cannot be guaranteed since the cryptographic proofs and analysis would no longer apply.)

For v1 encryption policies, the KDF only supports deriving per-file encryption keys. It works by encrypting the master key with AES-128-ECB, using the file's 16-byte nonce as the AES key. The resulting ciphertext is used as the derived key. If the ciphertext is longer than needed, then it is truncated to the needed length.

For v2 encryption policies, the KDF is HKDF-SHA512. The master key is passed as the ”input keying material”, no salt is used, and a distinct ”application-specific information string” is used for each distinct key to be derived. For example, when a per-file encryption key is derived, the application-specific information string is the file's nonce prefixed with ”fscrypt\0” and a context byte. Different context bytes are used for other types of derived keys.

HKDF-SHA512 is preferred to the original AES-128-ECB based KDF because HKDF is more flexible, is nonreversible, and evenly distributes entropy from the master key. HKDF is also standardized and widely used by other software, whereas the AES-128-ECB based KDF is ad-hoc.

### Per-file encryption keys

Since each master key can protect many files, it is necessary to ”tweak” the encryption of each file so that the same plaintext in two files doesn't map to the same ciphertext, or vice versa. In most cases, fscrypt does this by deriving per-file keys. When a new encrypted inode (regular file, directory, or symlink) is created, fscrypt randomly generates a 16-byte nonce and stores it in the inode's encryption xattr. Then, it uses a KDF (as described in *Key derivation function*) to derive the file's key from the master key and nonce.

Key derivation was chosen over key wrapping because wrapped keys would require larger xattrs which would be less likely to fit in-line in the filesystem's inode table, and there didn't appear to be any significant advantages to key wrapping. In particular, currently there is no requirement to support unlocking a file with multiple alternative master keys or to support rotating master keys. Instead, the master keys may be wrapped in userspace, e.g. as is done by the fscrypt tool.

### DIRECT_KEY policies

The Adiantum encryption mode (see *Encryption modes and usage*) is suitable for both contents and filenames encryption, and it accepts long IVs --- long enough to hold both an 8-byte data unit index and a 16-byte per-file nonce. Also, the overhead of each Adiantum key is greater than that of an AES-256-XTS key.

Therefore, to improve performance and save memory, for Adiantum a "direct key" configuration is supported. When the user has enabled this by setting FSCRYPT_POLICY_FLAG_DIRECT_KEY in the fscrypt policy, per-file encryption keys are not used. Instead, whenever any data (contents or filenames) is encrypted, the file's 16-byte nonce is included in the IV. Moreover:

- For v1 encryption policies, the encryption is done directly with the master key. Because of this, users **must not** use the same master key for any other purpose, even for other v1 policies.
- For v2 encryption policies, the encryption is done with a per-mode key derived using the KDF. Users may use the same master key for other v2 encryption policies.

### IV_INO_LBLK_64 policies

When FSCRYPT_POLICY_FLAG_IV_INO_LBLK_64 is set in the fscrypt policy, the encryption keys are derived from the master key, encryption mode number, and filesystem UUID. This normally results in all files protected by the same master key sharing a single contents encryption key and a single filenames encryption key. To still encrypt different files' data differently, inode numbers are included in the IVs. Consequently, shrinking the filesystem may not be allowed.

This format is optimized for use with inline encryption hardware compliant with the UFS standard, which supports only 64 IV bits per I/O request and may have only a small number of keyslots.

### IV_INO_LBLK_32 policies

IV_INO_LBLK_32 policies work like IV_INO_LBLK_64, except that for IV_INO_LBLK_32, the inode number is hashed with SipHash-2-4 (where the SipHash key is derived from the master key) and added to the file data unit index mod 2^32 to produce a 32-bit IV.

This format is optimized for use with inline encryption hardware compliant with the eMMC v5.2 standard, which supports only 32 IV bits per I/O request and may have only a small number of keyslots. This format results in some level of IV reuse, so it should only be used when necessary due to hardware limitations.

#### Key identifiers

For master keys used for v2 encryption policies, a unique 16-byte "key identifier" is also derived using the KDF. This value is stored in the clear, since it is needed to reliably identify the key itself.

#### Dirhash keys

For directories that are indexed using a secret-keyed dirhash over the plaintext filenames, the KDF is also used to derive a 128-bit SipHash-2-4 key per directory in order to hash filenames. This works just like deriving a per-file encryption key, except that a different KDF context is used. Currently, only casefolded ("case-insensitive") encrypted directories use this style of hashing.

### 2.2.4 Encryption modes and usage

fscrypt allows one encryption mode to be specified for file contents and one encryption mode to be specified for filenames. Different directory trees are permitted to use different encryption modes.

#### Supported modes

Currently, the following pairs of encryption modes are supported:

- AES-256-XTS for contents and AES-256-CTS-CBC for filenames
- AES-256-XTS for contents and AES-256-HCTR2 for filenames
- Adiantum for both contents and filenames
- AES-128-CBC-ESSIV for contents and AES-128-CTS-CBC for filenames
- SM4-XTS for contents and SM4-CTS-CBC for filenames

Authenticated encryption modes are not currently supported because of the difficulty of dealing with ciphertext expansion. Therefore, contents encryption uses a block cipher in XTS mode or CBC-ESSIV mode, or a wide-block cipher. Filenames encryption uses a block cipher in CTS-CBC mode or a wide-block cipher.

The (AES-256-XTS, AES-256-CTS-CBC) pair is the recommended default. It is also the only option that is *guaranteed* to always be supported if the kernel supports fscrypt at all; see *Kernel config options*.

The (AES-256-XTS, AES-256-HCTR2) pair is also a good choice that upgrades the filenames encryption to use a wide-block cipher. (A *wide-block cipher*, also called a tweakable super-pseudorandom permutation, has the property that changing one bit scrambles the entire result.) As described in *Filenames encryption*, a wide-block cipher is the ideal mode for the problem domain, though CTS-CBC is the "least bad" choice among the alternatives. For more information about HCTR2, see the HCTR2 paper.

Adiantum is recommended on systems where AES is too slow due to lack of hardware acceleration for AES. Adiantum is a wide-block cipher that uses XChaCha12 and AES-256 as its underlying components. Most of the work is done by XChaCha12, which is much faster than AES

when AES acceleration is unavailable. For more information about Adiantum, see the Adiantum paper.

The (AES-128-CBC-ESSIV, AES-128-CTS-CBC) pair exists only to support systems whose only form of AES acceleration is an off-CPU crypto accelerator such as CAAM or CESA that does not support XTS.

The remaining mode pairs are the ”national pride ciphers”:

- (SM4-XTS, SM4-CTS-CBC)

Generally speaking, these ciphers aren't ”bad” per se, but they receive limited security review compared to the usual choices such as AES and ChaCha. They also don't bring much new to the table. It is suggested to only use these ciphers where their use is mandated.

### Kernel config options

Enabling fscrypt support (CONFIG_FS_ENCRYPTION) automatically pulls in only the basic support from the crypto API needed to use AES-256-XTS and AES-256-CTS-CBC encryption. For optimal performance, it is strongly recommended to also enable any available platform-specific kconfig options that provide acceleration for the algorithm(s) you wish to use. Support for any ”non-default” encryption modes typically requires extra kconfig options as well.

Below, some relevant options are listed by encryption mode. Note, acceleration options not listed below may be available for your platform; refer to the kconfig menus. File contents encryption can also be configured to use inline encryption hardware instead of the kernel crypto API (see *Inline encryption support*); in that case, the file contents mode doesn't need to supported in the kernel crypto API, but the filenames mode still does.

- **AES-256-XTS and AES-256-CTS-CBC**
  - **Recommended:**
    - arm64: CONFIG_CRYPTO_AES_ARM64_CE_BLK
    - x86: CONFIG_CRYPTO_AES_NI_INTEL
- **AES-256-HCTR2**
  - **Mandatory:**
    - CONFIG_CRYPTO_HCTR2
  - **Recommended:**
    - arm64: CONFIG_CRYPTO_AES_ARM64_CE_BLK
    - arm64: CONFIG_CRYPTO_POLYVAL_ARM64_CE
    - x86: CONFIG_CRYPTO_AES_NI_INTEL
    - x86: CONFIG_CRYPTO_POLYVAL_CLMUL_NI
- **Adiantum**
  - **Mandatory:**
    - CONFIG_CRYPTO_ADIANTUM
  - **Recommended:**
    - arm32: CONFIG_CRYPTO_CHACHA20_NEON

        * arm32: CONFIG_CRYPTO_NHPOLY1305_NEON
        * arm64: CONFIG_CRYPTO_CHACHA20_NEON
        * arm64: CONFIG_CRYPTO_NHPOLY1305_NEON
        * x86: CONFIG_CRYPTO_CHACHA20_X86_64
        * x86: CONFIG_CRYPTO_NHPOLY1305_SSE2
        * x86: CONFIG_CRYPTO_NHPOLY1305_AVX2

- **AES-128-CBC-ESSIV and AES-128-CTS-CBC:**
    - **Mandatory:**
        * CONFIG_CRYPTO_ESSIV
        * CONFIG_CRYPTO_SHA256 or another SHA-256 implementation
    - **Recommended:**
        * AES-CBC acceleration

fscrypt also uses HMAC-SHA512 for key derivation, so enabling SHA-512 acceleration is recommended:

- **SHA-512**
    - **Recommended:**
        * arm64: CONFIG_CRYPTO_SHA512_ARM64_CE
        * x86: CONFIG_CRYPTO_SHA512_SSSE3

#### Contents encryption

For contents encryption, each file's contents is divided into ”data units”. Each data unit is encrypted independently. The IV for each data unit incorporates the zero-based index of the data unit within the file. This ensures that each data unit within a file is encrypted differently, which is essential to prevent leaking information.

Note: the encryption depending on the offset into the file means that operations like ”collapse range” and ”insert range” that rearrange the extent mapping of files are not supported on encrypted files.

There are two cases for the sizes of the data units:

- Fixed-size data units. This is how all filesystems other than UBIFS work. A file's data units are all the same size; the last data unit is zero-padded if needed. By default, the data unit size is equal to the filesystem block size. On some filesystems, users can select a sub-block data unit size via the `log2_data_unit_size` field of the encryption policy; see *FS_IOC_SET_ENCRYPTION_POLICY*.
- Variable-size data units. This is what UBIFS does. Each ”UBIFS data node” is treated as a crypto data unit. Each contains variable length, possibly compressed data, zero-padded to the next 16-byte boundary. Users cannot select a sub-block data unit size on UBIFS.

In the case of compression + encryption, the compressed data is encrypted. UBIFS compression works as described above. f2fs compression works a bit differently; it compresses a number of

filesystem blocks into a smaller number of filesystem blocks. Therefore a f2fs-compressed file still uses fixed-size data units, and it is encrypted in a similar way to a file containing holes.

As mentioned in *Key hierarchy*, the default encryption setting uses per-file keys. In this case, the IV for each data unit is simply the index of the data unit in the file. However, users can select an encryption setting that does not use per-file keys. For these, some kind of file identifier is incorporated into the IVs as follows:

- With *DIRECT_KEY policies*, the data unit index is placed in bits 0-63 of the IV, and the file's nonce is placed in bits 64-191.
- With *IV_INO_LBLK_64 policies*, the data unit index is placed in bits 0-31 of the IV, and the file's inode number is placed in bits 32-63. This setting is only allowed when data unit indices and inode numbers fit in 32 bits.
- With *IV_INO_LBLK_32 policies*, the file's inode number is hashed and added to the data unit index. The resulting value is truncated to 32 bits and placed in bits 0-31 of the IV. This setting is only allowed when data unit indices and inode numbers fit in 32 bits.

The byte order of the IV is always little endian.

If the user selects FSCRYPT_MODE_AES_128_CBC for the contents mode, an ESSIV layer is automatically included. In this case, before the IV is passed to AES-128-CBC, it is encrypted with AES-256 where the AES-256 key is the SHA-256 hash of the file's contents encryption key.

### Filenames encryption

For filenames, each full filename is encrypted at once. Because of the requirements to retain support for efficient directory lookups and filenames of up to 255 bytes, the same IV is used for every filename in a directory.

However, each encrypted directory still uses a unique key, or alternatively has the file's nonce (for *DIRECT_KEY policies*) or inode number (for *IV_INO_LBLK_64 policies*) included in the IVs. Thus, IV reuse is limited to within a single directory.

With CTS-CBC, the IV reuse means that when the plaintext filenames share a common prefix at least as long as the cipher block size (16 bytes for AES), the corresponding encrypted filenames will also share a common prefix. This is undesirable. Adiantum and HCTR2 do not have this weakness, as they are wide-block encryption modes.

All supported filenames encryption modes accept any plaintext length >= 16 bytes; cipher block alignment is not required. However, filenames shorter than 16 bytes are NUL-padded to 16 bytes before being encrypted. In addition, to reduce leakage of filename lengths via their ciphertexts, all filenames are NUL-padded to the next 4, 8, 16, or 32-byte boundary (configurable). 32 is recommended since this provides the best confidentiality, at the cost of making directory entries consume slightly more space. Note that since NUL (`\0`) is not otherwise a valid character in filenames, the padding will never produce duplicate plaintexts.

Symbolic link targets are considered a type of filename and are encrypted in the same way as filenames in directory entries, except that IV reuse is not a problem as each symlink has its own inode.

### 2.2.5 User API

#### Setting an encryption policy

#### FS_IOC_SET_ENCRYPTION_POLICY

The FS_IOC_SET_ENCRYPTION_POLICY ioctl sets an encryption policy on an empty directory or verifies that a directory or regular file already has the specified encryption policy. It takes in a pointer to struct fscrypt_policy_v1 or struct fscrypt_policy_v2, defined as follows:

```
#define FSCRYPT_POLICY_V1               0
#define FSCRYPT_KEY_DESCRIPTOR_SIZE     8
struct fscrypt_policy_v1 {
        __u8 version;
        __u8 contents_encryption_mode;
        __u8 filenames_encryption_mode;
        __u8 flags;
        __u8 master_key_descriptor[FSCRYPT_KEY_DESCRIPTOR_SIZE];
};
#define fscrypt_policy  fscrypt_policy_v1

#define FSCRYPT_POLICY_V2               2
#define FSCRYPT_KEY_IDENTIFIER_SIZE     16
struct fscrypt_policy_v2 {
        __u8 version;
        __u8 contents_encryption_mode;
        __u8 filenames_encryption_mode;
        __u8 flags;
        __u8 log2_data_unit_size;
        __u8 __reserved[3];
        __u8 master_key_identifier[FSCRYPT_KEY_IDENTIFIER_SIZE];
};
```

This structure must be initialized as follows:

- `version` must be FSCRYPT_POLICY_V1 (0) if struct fscrypt_policy_v1 is used or FSCRYPT_POLICY_V2 (2) if struct fscrypt_policy_v2 is used. (Note: we refer to the original policy version as "v1", though its version code is really 0.) For new encrypted directories, use v2 policies.

- `contents_encryption_mode` and `filenames_encryption_mode` must be set to constants from `<linux/fscrypt.h>` which identify the encryption modes to use. If unsure, use FSCRYPT_MODE_AES_256_XTS (1) for `contents_encryption_mode` and FSCRYPT_MODE_AES_256_CTS (4) for `filenames_encryption_mode`. For details, see *Encryption modes and usage*.

  v1 encryption policies only support three combinations of modes: (FSCRYPT_MODE_AES_256_XTS, FSCRYPT_MODE_AES_256_CTS), (FSCRYPT_MODE_AES_128_CBC, FSCRYPT_MODE_AES_128_CTS), and (FSCRYPT_MODE_ADIANTUM, FSCRYPT_MODE_ADIANTUM). v2 policies support all combinations documented in *Supported modes*.

- `flags` contains optional flags from `<linux/fscrypt.h>`:

  - FSCRYPT_POLICY_FLAGS_PAD_*: The amount of NUL padding to use when encrypting filenames. If unsure, use FSCRYPT_POLICY_FLAGS_PAD_32 (0x3).
  - FSCRYPT_POLICY_FLAG_DIRECT_KEY: See *DIRECT_KEY policies*.
  - FSCRYPT_POLICY_FLAG_IV_INO_LBLK_64: See *IV_INO_LBLK_64 policies*.
  - FSCRYPT_POLICY_FLAG_IV_INO_LBLK_32: See *IV_INO_LBLK_32 policies*.

  v1 encryption policies only support the PAD_* and DIRECT_KEY flags. The other flags are only supported by v2 encryption policies.

  The DIRECT_KEY, IV_INO_LBLK_64, and IV_INO_LBLK_32 flags are mutually exclusive.

- `log2_data_unit_size` is the log2 of the data unit size in bytes, or 0 to select the default data unit size. The data unit size is the granularity of file contents encryption. For example, setting `log2_data_unit_size` to 12 causes file contents be passed to the underlying encryption algorithm (such as AES-256-XTS) in 4096-byte data units, each with its own IV.

  Not all filesystems support setting `log2_data_unit_size`. ext4 and f2fs support it since Linux v6.7. On filesystems that support it, the supported nonzero values are 9 through the log2 of the filesystem block size, inclusively. The default value of 0 selects the filesystem block size.

  The main use case for `log2_data_unit_size` is for selecting a data unit size smaller than the filesystem block size for compatibility with inline encryption hardware that only supports smaller data unit sizes. `/sys/block/$disk/queue/crypto/` may be useful for checking which data unit sizes are supported by a particular system's inline encryption hardware.

  Leave this field zeroed unless you are certain you need it. Using an unnecessarily small data unit size reduces performance.

- For v2 encryption policies, `__reserved` must be zeroed.

- For v1 encryption policies, `master_key_descriptor` specifies how to find the master key in a keyring; see *Adding keys*. It is up to userspace to choose a unique `master_key_descriptor` for each master key. The e4crypt and fscrypt tools use the first 8 bytes of `SHA-512(SHA-512(master_key))`, but this particular scheme is not required. Also, the master key need not be in the keyring yet when FS_IOC_SET_ENCRYPTION_POLICY is executed. However, it must be added before any files can be created in the encrypted directory.

  For v2 encryption policies, `master_key_descriptor` has been replaced with `master_key_identifier`, which is longer and cannot be arbitrarily chosen. Instead, the key must first be added using *FS_IOC_ADD_ENCRYPTION_KEY*. Then, the `key_spec.u.identifier` the kernel returned in the struct fscrypt_add_key_arg must be used as the `master_key_identifier` in struct fscrypt_policy_v2.

If the file is not yet encrypted, then FS_IOC_SET_ENCRYPTION_POLICY verifies that the file is an empty directory. If so, the specified encryption policy is assigned to the directory, turning it into an encrypted directory. After that, and after providing the corresponding master key as described in *Adding keys*, all regular files, directories (recursively), and symlinks created in the directory will be encrypted, inheriting the same encryption policy. The filenames in the directory's entries will be encrypted as well.

Alternatively, if the file is already encrypted, then FS_IOC_SET_ENCRYPTION_POLICY validates that the specified encryption policy exactly matches the actual one. If they match, then

the ioctl returns 0. Otherwise, it fails with EEXIST. This works on both regular files and directories, including nonempty directories.

When a v2 encryption policy is assigned to a directory, it is also required that either the specified key has been added by the current user or that the caller has CAP_FOWNER in the initial user namespace. (This is needed to prevent a user from encrypting their data with another user's key.) The key must remain added while FS_IOC_SET_ENCRYPTION_POLICY is executing. However, if the new encrypted directory does not need to be accessed immediately, then the key can be removed right away afterwards.

Note that the ext4 filesystem does not allow the root directory to be encrypted, even if it is empty. Users who want to encrypt an entire filesystem with one key should consider using dm-crypt instead.

FS_IOC_SET_ENCRYPTION_POLICY can fail with the following errors:

- `EACCES`: the file is not owned by the process's uid, nor does the process have the CAP_FOWNER capability in a namespace with the file owner's uid mapped
- `EEXIST`: the file is already encrypted with an encryption policy different from the one specified
- `EINVAL`: an invalid encryption policy was specified (invalid version, mode(s), or flags; or reserved bits were set); or a v1 encryption policy was specified but the directory has the casefold flag enabled (casefolding is incompatible with v1 policies).
- `ENOKEY`: a v2 encryption policy was specified, but the key with the specified `master_key_identifier` has not been added, nor does the process have the CAP_FOWNER capability in the initial user namespace
- `ENOTDIR`: the file is unencrypted and is a regular file, not a directory
- `ENOTEMPTY`: the file is unencrypted and is a nonempty directory
- `ENOTTY`: this type of filesystem does not implement encryption
- `EOPNOTSUPP`: the kernel was not configured with encryption support for filesystems, or the filesystem superblock has not had encryption enabled on it. (For example, to use encryption on an ext4 filesystem, CONFIG_FS_ENCRYPTION must be enabled in the kernel config, and the superblock must have had the "encrypt" feature flag enabled using `tune2fs -O encrypt` or `mkfs.ext4 -O encrypt`.)
- `EPERM`: this directory may not be encrypted, e.g. because it is the root directory of an ext4 filesystem
- `EROFS`: the filesystem is readonly

### Getting an encryption policy

Two ioctls are available to get a file's encryption policy:

- *FS_IOC_GET_ENCRYPTION_POLICY_EX*
- *FS_IOC_GET_ENCRYPTION_POLICY*

The extended (_EX) version of the ioctl is more general and is recommended to use when possible. However, on older kernels only the original ioctl is available. Applications should try the extended version, and if it fails with ENOTTY fall back to the original version.

### FS_IOC_GET_ENCRYPTION_POLICY_EX

The FS_IOC_GET_ENCRYPTION_POLICY_EX ioctl retrieves the encryption policy, if any, for a directory or regular file. No additional permissions are required beyond the ability to open the file. It takes in a pointer to struct fscrypt_get_policy_ex_arg, defined as follows:

```
struct fscrypt_get_policy_ex_arg {
        __u64 policy_size; /* input/output */
        union {
                __u8 version;
                struct fscrypt_policy_v1 v1;
                struct fscrypt_policy_v2 v2;
        } policy; /* output */
};
```

The caller must initialize `policy_size` to the size available for the policy struct, i.e. `sizeof(arg.policy)`.

On success, the policy struct is returned in `policy`, and its actual size is returned in `policy_size`. `policy.version` should be checked to determine the version of policy returned. Note that the version code for the "v1" policy is actually 0 (FSCRYPT_POLICY_V1).

FS_IOC_GET_ENCRYPTION_POLICY_EX can fail with the following errors:

- `EINVAL`: the file is encrypted, but it uses an unrecognized encryption policy version
- `ENODATA`: the file is not encrypted
- `ENOTTY`: this type of filesystem does not implement encryption, or this kernel is too old to support FS_IOC_GET_ENCRYPTION_POLICY_EX (try FS_IOC_GET_ENCRYPTION_POLICY instead)
- `EOPNOTSUPP`: the kernel was not configured with encryption support for this filesystem, or the filesystem superblock has not had encryption enabled on it
- `EOVERFLOW`: the file is encrypted and uses a recognized encryption policy version, but the policy struct does not fit into the provided buffer

Note: if you only need to know whether a file is encrypted or not, on most filesystems it is also possible to use the FS_IOC_GETFLAGS ioctl and check for FS_ENCRYPT_FL, or to use the statx() system call and check for STATX_ATTR_ENCRYPTED in stx_attributes.

### FS_IOC_GET_ENCRYPTION_POLICY

The FS_IOC_GET_ENCRYPTION_POLICY ioctl can also retrieve the encryption policy, if any, for a directory or regular file. However, unlike *FS_IOC_GET_ENCRYPTION_POLICY_EX*, FS_IOC_GET_ENCRYPTION_POLICY only supports the original policy version. It takes in a pointer directly to struct fscrypt_policy_v1 rather than struct fscrypt_get_policy_ex_arg.

The error codes for FS_IOC_GET_ENCRYPTION_POLICY are the same as those for FS_IOC_GET_ENCRYPTION_POLICY_EX, except that FS_IOC_GET_ENCRYPTION_POLICY also returns `EINVAL` if the file is encrypted using a newer encryption policy version.

### Getting the per-filesystem salt

Some filesystems, such as ext4 and F2FS, also support the deprecated ioctl FS_IOC_GET_ENCRYPTION_PWSALT. This ioctl retrieves a randomly generated 16-byte value stored in the filesystem superblock. This value is intended to used as a salt when deriving an encryption key from a passphrase or other low-entropy user credential.

FS_IOC_GET_ENCRYPTION_PWSALT is deprecated. Instead, prefer to generate and manage any needed salt(s) in userspace.

### Getting a file's encryption nonce

Since Linux v5.7, the ioctl FS_IOC_GET_ENCRYPTION_NONCE is supported. On encrypted files and directories it gets the inode's 16-byte nonce. On unencrypted files and directories, it fails with ENODATA.

This ioctl can be useful for automated tests which verify that the encryption is being done correctly. It is not needed for normal use of fscrypt.

### Adding keys

#### FS_IOC_ADD_ENCRYPTION_KEY

The FS_IOC_ADD_ENCRYPTION_KEY ioctl adds a master encryption key to the filesystem, making all files on the filesystem which were encrypted using that key appear "unlocked", i.e. in plaintext form. It can be executed on any file or directory on the target filesystem, but using the filesystem's root directory is recommended. It takes in a pointer to struct fscrypt_add_key_arg, defined as follows:

```
struct fscrypt_add_key_arg {
        struct fscrypt_key_specifier key_spec;
        __u32 raw_size;
        __u32 key_id;
        __u32 __reserved[8];
        __u8 raw[];
};

#define FSCRYPT_KEY_SPEC_TYPE_DESCRIPTOR        1
#define FSCRYPT_KEY_SPEC_TYPE_IDENTIFIER        2

struct fscrypt_key_specifier {
        __u32 type;     /* one of FSCRYPT_KEY_SPEC_TYPE_* */
        __u32 __reserved;
        union {
                __u8 __reserved[32]; /* reserve some extra space */
                __u8 descriptor[FSCRYPT_KEY_DESCRIPTOR_SIZE];
                __u8 identifier[FSCRYPT_KEY_IDENTIFIER_SIZE];
        } u;
};

struct fscrypt_provisioning_key_payload {
```

```
        __u32 type;
        __u32 __reserved;
        __u8 raw[];
};
```

struct fscrypt_add_key_arg must be zeroed, then initialized as follows:

- If the key is being added for use by v1 encryption policies, then `key_spec.type` must contain FSCRYPT_KEY_SPEC_TYPE_DESCRIPTOR, and `key_spec.u.descriptor` must contain the descriptor of the key being added, corresponding to the value in the `master_key_descriptor` field of struct fscrypt_policy_v1. To add this type of key, the calling process must have the CAP_SYS_ADMIN capability in the initial user namespace.

  Alternatively, if the key is being added for use by v2 encryption policies, then `key_spec.type` must contain FSCRYPT_KEY_SPEC_TYPE_IDENTIFIER, and `key_spec.u.identifier` is an *output* field which the kernel fills in with a cryptographic hash of the key. To add this type of key, the calling process does not need any privileges. However, the number of keys that can be added is limited by the user's quota for the keyrings service (see `Documentation/security/keys/core.rst`).

- `raw_size` must be the size of the `raw` key provided, in bytes. Alternatively, if `key_id` is nonzero, this field must be 0, since in that case the size is implied by the specified Linux keyring key.

- `key_id` is 0 if the raw key is given directly in the `raw` field. Otherwise `key_id` is the ID of a Linux keyring key of type "fscrypt-provisioning" whose payload is struct fscrypt_provisioning_key_payload whose `raw` field contains the raw key and whose `type` field matches `key_spec.type`. Since `raw` is variable-length, the total size of this key's payload must be `sizeof(struct fscrypt_provisioning_key_payload)` plus the raw key size. The process must have Search permission on this key.

  Most users should leave this 0 and specify the raw key directly. The support for specifying a Linux keyring key is intended mainly to allow re-adding keys after a filesystem is unmounted and re-mounted, without having to store the raw keys in userspace memory.

- `raw` is a variable-length field which must contain the actual key, `raw_size` bytes long. Alternatively, if `key_id` is nonzero, then this field is unused.

For v2 policy keys, the kernel keeps track of which user (identified by effective user ID) added the key, and only allows the key to be removed by that user --- or by "root", if they use *FS_IOC_REMOVE_ENCRYPTION_KEY_ALL_USERS*.

However, if another user has added the key, it may be desirable to prevent that other user from unexpectedly removing it. Therefore, FS_IOC_ADD_ENCRYPTION_KEY may also be used to add a v2 policy key *again*, even if it's already added by other user(s). In this case, FS_IOC_ADD_ENCRYPTION_KEY will just install a claim to the key for the current user, rather than actually add the key again (but the raw key must still be provided, as a proof of knowledge).

FS_IOC_ADD_ENCRYPTION_KEY returns 0 if either the key or a claim to the key was either added or already exists.

FS_IOC_ADD_ENCRYPTION_KEY can fail with the following errors:

- `EACCES`: FSCRYPT_KEY_SPEC_TYPE_DESCRIPTOR was specified, but the caller does not have the CAP_SYS_ADMIN capability in the initial user namespace; or the raw key was specified by Linux key ID but the process lacks Search permission on the key.

- EDQUOT: the key quota for this user would be exceeded by adding the key
- EINVAL: invalid key size or key specifier type, or reserved bits were set
- EKEYREJECTED: the raw key was specified by Linux key ID, but the key has the wrong type
- ENOKEY: the raw key was specified by Linux key ID, but no key exists with that ID
- ENOTTY: this type of filesystem does not implement encryption
- EOPNOTSUPP: the kernel was not configured with encryption support for this filesystem, or the filesystem superblock has not had encryption enabled on it

### Legacy method

For v1 encryption policies, a master encryption key can also be provided by adding it to a process-subscribed keyring, e.g. to a session keyring, or to a user keyring if the user keyring is linked into the session keyring.

This method is deprecated (and not supported for v2 encryption policies) for several reasons. First, it cannot be used in combination with FS_IOC_REMOVE_ENCRYPTION_KEY (see *Removing keys*), so for removing a key a workaround such as keyctl_unlink() in combination with `sync; echo 2 > /proc/sys/vm/drop_caches` would have to be used. Second, it doesn't match the fact that the locked/unlocked status of encrypted files (i.e. whether they appear to be in plaintext form or in ciphertext form) is global. This mismatch has caused much confusion as well as real problems when processes running under different UIDs, such as a `sudo` command, need to access encrypted files.

Nevertheless, to add a key to one of the process-subscribed keyrings, the add_key() system call can be used (see: `Documentation/security/keys/core.rst`). The key type must be "logon"; keys of this type are kept in kernel memory and cannot be read back by userspace. The key description must be "fscrypt:" followed by the 16-character lower case hex representation of the `master_key_descriptor` that was set in the encryption policy. The key payload must conform to the following structure:

```
#define FSCRYPT_MAX_KEY_SIZE            64

struct fscrypt_key {
        __u32 mode;
        __u8 raw[FSCRYPT_MAX_KEY_SIZE];
        __u32 size;
};
```

`mode` is ignored; just set it to 0. The actual key is provided in `raw` with `size` indicating its size in bytes. That is, the bytes `raw[0..size-1]` (inclusive) are the actual key.

The key description prefix "fscrypt:" may alternatively be replaced with a filesystem-specific prefix such as "ext4:". However, the filesystem-specific prefixes are deprecated and should not be used in new programs.

### Removing keys

Two ioctls are available for removing a key that was added by *FS_IOC_ADD_ENCRYPTION_KEY*:

- *FS_IOC_REMOVE_ENCRYPTION_KEY*
- *FS_IOC_REMOVE_ENCRYPTION_KEY_ALL_USERS*

These two ioctls differ only in cases where v2 policy keys are added or removed by non-root users.

These ioctls don't work on keys that were added via the legacy process-subscribed keyrings mechanism.

Before using these ioctls, read the *Kernel memory compromise* section for a discussion of the security goals and limitations of these ioctls.

### FS_IOC_REMOVE_ENCRYPTION_KEY

The FS_IOC_REMOVE_ENCRYPTION_KEY ioctl removes a claim to a master encryption key from the filesystem, and possibly removes the key itself. It can be executed on any file or directory on the target filesystem, but using the filesystem's root directory is recommended. It takes in a pointer to struct fscrypt_remove_key_arg, defined as follows:

```
struct fscrypt_remove_key_arg {
        struct fscrypt_key_specifier key_spec;
#define FSCRYPT_KEY_REMOVAL_STATUS_FLAG_FILES_BUSY      0x00000001
#define FSCRYPT_KEY_REMOVAL_STATUS_FLAG_OTHER_USERS     0x00000002
        __u32 removal_status_flags;     /* output */
        __u32 __reserved[5];
};
```

This structure must be zeroed, then initialized as follows:

- The key to remove is specified by `key_spec`:
  - To remove a key used by v1 encryption policies, set `key_spec.type` to FSCRYPT_KEY_SPEC_TYPE_DESCRIPTOR and fill in `key_spec.u.descriptor`. To remove this type of key, the calling process must have the CAP_SYS_ADMIN capability in the initial user namespace.
  - To remove a key used by v2 encryption policies, set `key_spec.type` to FSCRYPT_KEY_SPEC_TYPE_IDENTIFIER and fill in `key_spec.u.identifier`.

For v2 policy keys, this ioctl is usable by non-root users. However, to make this possible, it actually just removes the current user's claim to the key, undoing a single call to FS_IOC_ADD_ENCRYPTION_KEY. Only after all claims are removed is the key really removed.

For example, if FS_IOC_ADD_ENCRYPTION_KEY was called with uid 1000, then the key will be ”claimed” by uid 1000, and FS_IOC_REMOVE_ENCRYPTION_KEY will only succeed as uid 1000. Or, if both uids 1000 and 2000 added the key, then for each uid FS_IOC_REMOVE_ENCRYPTION_KEY will only remove their own claim. Only once *both* are removed is the key really removed. (Think of it like unlinking a file that may have hard links.)

If FS_IOC_REMOVE_ENCRYPTION_KEY really removes the key, it will also try to ”lock” all files that had been unlocked with the key. It won't lock files that are still in-use, so this ioctl is

expected to be used in cooperation with userspace ensuring that none of the files are still open. However, if necessary, this ioctl can be executed again later to retry locking any remaining files.

FS_IOC_REMOVE_ENCRYPTION_KEY returns 0 if either the key was removed (but may still have files remaining to be locked), the user's claim to the key was removed, or the key was already removed but had files remaining to be the locked so the ioctl retried locking them. In any of these cases, `removal_status_flags` is filled in with the following informational status flags:

- `FSCRYPT_KEY_REMOVAL_STATUS_FLAG_FILES_BUSY`: set if some file(s) are still in-use. Not guaranteed to be set in the case where only the user's claim to the key was removed.
- `FSCRYPT_KEY_REMOVAL_STATUS_FLAG_OTHER_USERS`: set if only the user's claim to the key was removed, not the key itself

FS_IOC_REMOVE_ENCRYPTION_KEY can fail with the following errors:

- `EACCES`: The FSCRYPT_KEY_SPEC_TYPE_DESCRIPTOR key specifier type was specified, but the caller does not have the CAP_SYS_ADMIN capability in the initial user namespace
- `EINVAL`: invalid key specifier type, or reserved bits were set
- `ENOKEY`: the key object was not found at all, i.e. it was never added in the first place or was already fully removed including all files locked; or, the user does not have a claim to the key (but someone else does).
- `ENOTTY`: this type of filesystem does not implement encryption
- `EOPNOTSUPP`: the kernel was not configured with encryption support for this filesystem, or the filesystem superblock has not had encryption enabled on it

#### FS_IOC_REMOVE_ENCRYPTION_KEY_ALL_USERS

FS_IOC_REMOVE_ENCRYPTION_KEY_ALL_USERS is exactly the same as *FS_IOC_REMOVE_ENCRYPTION_KEY*, except that for v2 policy keys, the ALL_USERS version of the ioctl will remove all users' claims to the key, not just the current user's. I.e., the key itself will always be removed, no matter how many users have added it. This difference is only meaningful if non-root users are adding and removing keys.

Because of this, FS_IOC_REMOVE_ENCRYPTION_KEY_ALL_USERS also requires "root", namely the CAP_SYS_ADMIN capability in the initial user namespace. Otherwise it will fail with EACCES.

### Getting key status

#### FS_IOC_GET_ENCRYPTION_KEY_STATUS

The FS_IOC_GET_ENCRYPTION_KEY_STATUS ioctl retrieves the status of a master encryption key. It can be executed on any file or directory on the target filesystem, but using the filesystem's root directory is recommended. It takes in a pointer to struct fscrypt_get_key_status_arg, defined as follows:

```
struct fscrypt_get_key_status_arg {
        /* input */
```

```
        struct fscrypt_key_specifier key_spec;
        __u32 __reserved[6];

        /* output */
#define FSCRYPT_KEY_STATUS_ABSENT               1
#define FSCRYPT_KEY_STATUS_PRESENT              2
#define FSCRYPT_KEY_STATUS_INCOMPLETELY_REMOVED 3
        __u32 status;
#define FSCRYPT_KEY_STATUS_FLAG_ADDED_BY_SELF   0x00000001
        __u32 status_flags;
        __u32 user_count;
        __u32 __out_reserved[13];
};
```

The caller must zero all input fields, then fill in `key_spec`:

- To get the status of a key for v1 encryption policies, set `key_spec.type` to FSCRYPT_KEY_SPEC_TYPE_DESCRIPTOR and fill in `key_spec.u.descriptor`.
- To get the status of a key for v2 encryption policies, set `key_spec.type` to FSCRYPT_KEY_SPEC_TYPE_IDENTIFIER and fill in `key_spec.u.identifier`.

On success, 0 is returned and the kernel fills in the output fields:

- `status` indicates whether the key is absent, present, or incompletely removed. Incompletely removed means that removal has been initiated, but some files are still in use; i.e., *FS_IOC_REMOVE_ENCRYPTION_KEY* returned 0 but set the informational status flag FSCRYPT_KEY_REMOVAL_STATUS_FLAG_FILES_BUSY.
- `status_flags` can contain the following flags:
  - `FSCRYPT_KEY_STATUS_FLAG_ADDED_BY_SELF` indicates that the key has added by the current user. This is only set for keys identified by `identifier` rather than by `descriptor`.
- `user_count` specifies the number of users who have added the key. This is only set for keys identified by `identifier` rather than by `descriptor`.

FS_IOC_GET_ENCRYPTION_KEY_STATUS can fail with the following errors:

- `EINVAL`: invalid key specifier type, or reserved bits were set
- `ENOTTY`: this type of filesystem does not implement encryption
- `EOPNOTSUPP`: the kernel was not configured with encryption support for this filesystem, or the filesystem superblock has not had encryption enabled on it

Among other use cases, FS_IOC_GET_ENCRYPTION_KEY_STATUS can be useful for determining whether the key for a given encrypted directory needs to be added before prompting the user for the passphrase needed to derive the key.

FS_IOC_GET_ENCRYPTION_KEY_STATUS can only get the status of keys in the filesystem-level keyring, i.e. the keyring managed by *FS_IOC_ADD_ENCRYPTION_KEY* and *FS_IOC_REMOVE_ENCRYPTION_KEY*. It cannot get the status of a key that has only been added for use by v1 encryption policies using the legacy mechanism involving process-subscribed keyrings.

### 2.2.6 Access semantics

#### With the key

With the encryption key, encrypted regular files, directories, and symlinks behave very similarly to their unencrypted counterparts --- after all, the encryption is intended to be transparent. However, astute users may notice some differences in behavior:

- Unencrypted files, or files encrypted with a different encryption policy (i.e. different key, modes, or flags), cannot be renamed or linked into an encrypted directory; see *Encryption policy enforcement*. Attempts to do so will fail with EXDEV. However, encrypted files can be renamed within an encrypted directory, or into an unencrypted directory.

  Note: ”moving” an unencrypted file into an encrypted directory, e.g. with the *mv* program, is implemented in userspace by a copy followed by a delete. Be aware that the original unencrypted data may remain recoverable from free space on the disk; prefer to keep all files encrypted from the very beginning. The *shred* program may be used to overwrite the source files but isn't guaranteed to be effective on all filesystems and storage devices.

- Direct I/O is supported on encrypted files only under some circumstances. For details, see *Direct I/O support*.

- The fallocate operations FALLOC_FL_COLLAPSE_RANGE and FALLOC_FL_INSERT_RANGE are not supported on encrypted files and will fail with EOPNOTSUPP.

- Online defragmentation of encrypted files is not supported. The EXT4_IOC_MOVE_EXT and F2FS_IOC_MOVE_RANGE ioctls will fail with EOPNOTSUPP.

- The ext4 filesystem does not support data journaling with encrypted regular files. It will fall back to ordered data mode instead.

- DAX (Direct Access) is not supported on encrypted files.

- The maximum length of an encrypted symlink is 2 bytes shorter than the maximum length of an unencrypted symlink. For example, on an EXT4 filesystem with a 4K block size, unencrypted symlinks can be up to 4095 bytes long, while encrypted symlinks can only be up to 4093 bytes long (both lengths excluding the terminating null).

Note that mmap *is* supported. This is possible because the pagecache for an encrypted file contains the plaintext, not the ciphertext.

#### Without the key

Some filesystem operations may be performed on encrypted regular files, directories, and symlinks even before their encryption key has been added, or after their encryption key has been removed:

- File metadata may be read, e.g. using stat().

- Directories may be listed, in which case the filenames will be listed in an encoded form derived from their ciphertext. The current encoding algorithm is described in *Filename hashing and encoding*. The algorithm is subject to change, but it is guaranteed that the presented filenames will be no longer than NAME_MAX bytes, will not contain the `/` or `\0` characters, and will uniquely identify directory entries.

The `.` and `..` directory entries are special. They are always present and are not encrypted or encoded.

- Files may be deleted. That is, nondirectory files may be deleted with unlink() as usual, and empty directories may be deleted with rmdir() as usual. Therefore, `rm` and `rm -r` will work as expected.
- Symlink targets may be read and followed, but they will be presented in encrypted form, similar to filenames in directories. Hence, they are unlikely to point to anywhere useful.

Without the key, regular files cannot be opened or truncated. Attempts to do so will fail with ENOKEY. This implies that any regular file operations that require a file descriptor, such as read(), write(), mmap(), fallocate(), and ioctl(), are also forbidden.

Also without the key, files of any type (including directories) cannot be created or linked into an encrypted directory, nor can a name in an encrypted directory be the source or target of a rename, nor can an O_TMPFILE temporary file be created in an encrypted directory. All such operations will fail with ENOKEY.

It is not currently possible to backup and restore encrypted files without the encryption key. This would require special APIs which have not yet been implemented.

### 2.2.7 Encryption policy enforcement

After an encryption policy has been set on a directory, all regular files, directories, and symbolic links created in that directory (recursively) will inherit that encryption policy. Special files --- that is, named pipes, device nodes, and UNIX domain sockets --- will not be encrypted.

Except for those special files, it is forbidden to have unencrypted files, or files encrypted with a different encryption policy, in an encrypted directory tree. Attempts to link or rename such a file into an encrypted directory will fail with EXDEV. This is also enforced during ->lookup() to provide limited protection against offline attacks that try to disable or downgrade encryption in known locations where applications may later write sensitive data. It is recommended that systems implementing a form of ”verified boot” take advantage of this by validating all top-level encryption policies prior to access.

### 2.2.8 Inline encryption support

By default, fscrypt uses the kernel crypto API for all cryptographic operations (other than HKDF, which fscrypt partially implements itself). The kernel crypto API supports hardware crypto accelerators, but only ones that work in the traditional way where all inputs and outputs (e.g. plaintexts and ciphertexts) are in memory. fscrypt can take advantage of such hardware, but the traditional acceleration model isn't particularly efficient and fscrypt hasn't been optimized for it.

Instead, many newer systems (especially mobile SoCs) have *inline encryption hardware* that can encrypt/decrypt data while it is on its way to/from the storage device. Linux supports inline encryption through a set of extensions to the block layer called *blk-crypto*. blk-crypto allows filesystems to attach encryption contexts to bios (I/O requests) to specify how the data will be encrypted or decrypted in-line. For more information about blk-crypto, see Documentation/block/inline-encryption.rst.

On supported filesystems (currently ext4 and f2fs), fscrypt can use blk-crypto instead of the kernel crypto API to encrypt/decrypt file contents. To enable this, set CON-

FIG_FS_ENCRYPTION_INLINE_CRYPT=y in the kernel configuration, and specify the ”inlinecrypt” mount option when mounting the filesystem.

Note that the ”inlinecrypt” mount option just specifies to use inline encryption when possible; it doesn't force its use. fscrypt will still fall back to using the kernel crypto API on files where the inline encryption hardware doesn't have the needed crypto capabilities (e.g. support for the needed encryption algorithm and data unit size) and where blk-crypto-fallback is unusable. (For blk-crypto-fallback to be usable, it must be enabled in the kernel configuration with CONFIG_BLK_INLINE_ENCRYPTION_FALLBACK=y.)

Currently fscrypt always uses the filesystem block size (which is usually 4096 bytes) as the data unit size. Therefore, it can only use inline encryption hardware that supports that data unit size.

Inline encryption doesn't affect the ciphertext or other aspects of the on-disk format, so users may freely switch back and forth between using ”inlinecrypt” and not using ”inlinecrypt”.

### 2.2.9 Direct I/O support

For direct I/O on an encrypted file to work, the following conditions must be met (in addition to the conditions for direct I/O on an unencrypted file):

- The file must be using inline encryption. Usually this means that the filesystem must be mounted with `-o inlinecrypt` and inline encryption hardware must be present. However, a software fallback is also available. For details, see *Inline encryption support*.
- The I/O request must be fully aligned to the filesystem block size. This means that the file position the I/O is targeting, the lengths of all I/O segments, and the memory addresses of all I/O buffers must be multiples of this value. Note that the filesystem block size may be greater than the logical block size of the block device.

If either of the above conditions is not met, then direct I/O on the encrypted file will fall back to buffered I/O.

### 2.2.10 Implementation details

#### Encryption context

An encryption policy is represented on-disk by struct fscrypt_context_v1 or struct fscrypt_context_v2. It is up to individual filesystems to decide where to store it, but normally it would be stored in a hidden extended attribute. It should *not* be exposed by the xattr-related system calls such as getxattr() and setxattr() because of the special semantics of the encryption xattr. (In particular, there would be much confusion if an encryption policy were to be added to or removed from anything other than an empty directory.) These structs are defined as follows:

```
#define FSCRYPT_FILE_NONCE_SIZE 16

#define FSCRYPT_KEY_DESCRIPTOR_SIZE  8
struct fscrypt_context_v1 {
        u8 version;
        u8 contents_encryption_mode;
        u8 filenames_encryption_mode;
        u8 flags;
        u8 master_key_descriptor[FSCRYPT_KEY_DESCRIPTOR_SIZE];
```

```
        u8 nonce[FSCRYPT_FILE_NONCE_SIZE];
};

#define FSCRYPT_KEY_IDENTIFIER_SIZE  16
struct fscrypt_context_v2 {
        u8 version;
        u8 contents_encryption_mode;
        u8 filenames_encryption_mode;
        u8 flags;
        u8 log2_data_unit_size;
        u8 __reserved[3];
        u8 master_key_identifier[FSCRYPT_KEY_IDENTIFIER_SIZE];
        u8 nonce[FSCRYPT_FILE_NONCE_SIZE];
};
```

The context structs contain the same information as the corresponding policy structs (see *Setting an encryption policy*), except that the context structs also contain a nonce. The nonce is randomly generated by the kernel and is used as KDF input or as a tweak to cause different files to be encrypted differently; see *Per-file encryption keys* and *DIRECT_KEY policies*.

### Data path changes

When inline encryption is used, filesystems just need to associate encryption contexts with bios to specify how the block layer or the inline encryption hardware will encrypt/decrypt the file contents.

When inline encryption isn't used, filesystems must encrypt/decrypt the file contents themselves, as described below:

For the read path (->read_folio()) of regular files, filesystems can read the ciphertext into the page cache and decrypt it in-place. The folio lock must be held until decryption has finished, to prevent the folio from becoming visible to userspace prematurely.

For the write path (->writepage()) of regular files, filesystems cannot encrypt data in-place in the page cache, since the cached plaintext must be preserved. Instead, filesystems must encrypt into a temporary buffer or "bounce page", then write out the temporary buffer. Some filesystems, such as UBIFS, already use temporary buffers regardless of encryption. Other filesystems, such as ext4 and F2FS, have to allocate bounce pages specially for encryption.

### Filename hashing and encoding

Modern filesystems accelerate directory lookups by using indexed directories. An indexed directory is organized as a tree keyed by filename hashes. When a ->lookup() is requested, the filesystem normally hashes the filename being looked up so that it can quickly find the corresponding directory entry, if any.

With encryption, lookups must be supported and efficient both with and without the encryption key. Clearly, it would not work to hash the plaintext filenames, since the plaintext filenames are unavailable without the key. (Hashing the plaintext filenames would also make it impossible for the filesystem's fsck tool to optimize encrypted directories.) Instead, filesystems hash the ciphertext filenames, i.e. the bytes actually stored on-disk in the directory entries. When asked

to do a ->lookup() with the key, the filesystem just encrypts the user-supplied name to get the ciphertext.

Lookups without the key are more complicated. The raw ciphertext may contain the `\0` and `/` characters, which are illegal in filenames. Therefore, readdir() must base64url-encode the ciphertext for presentation. For most filenames, this works fine; on ->lookup(), the filesystem just base64url-decodes the user-supplied name to get back to the raw ciphertext.

However, for very long filenames, base64url encoding would cause the filename length to exceed NAME_MAX. To prevent this, readdir() actually presents long filenames in an abbreviated form which encodes a strong ”hash” of the ciphertext filename, along with the optional filesystem-specific hash(es) needed for directory lookups. This allows the filesystem to still, with a high degree of confidence, map the filename given in ->lookup() back to a particular directory entry that was previously listed by readdir(). See struct fscrypt_nokey_name in the source for more details.

Note that the precise way that filenames are presented to userspace without the key is subject to change in the future. It is only meant as a way to temporarily present valid filenames so that commands like `rm -r` work as expected on encrypted directories.

### 2.2.11 Tests

To test fscrypt, use xfstests, which is Linux's de facto standard filesystem test suite. First, run all the tests in the ”encrypt” group on the relevant filesystem(s). One can also run the tests with the 'inlinecrypt' mount option to test the implementation for inline encryption support. For example, to test ext4 and f2fs encryption using kvm-xfstests:

```
kvm-xfstests -c ext4,f2fs -g encrypt
kvm-xfstests -c ext4,f2fs -g encrypt -m inlinecrypt
```

UBIFS encryption can also be tested this way, but it should be done in a separate command, and it takes some time for kvm-xfstests to set up emulated UBI volumes:

```
kvm-xfstests -c ubifs -g encrypt
```

No tests should fail. However, tests that use non-default encryption modes (e.g. generic/549 and generic/550) will be skipped if the needed algorithms were not built into the kernel's crypto API. Also, tests that access the raw block device (e.g. generic/399, generic/548, generic/549, generic/550) will be skipped on UBIFS.

Besides running the ”encrypt” group tests, for ext4 and f2fs it's also possible to run most xfstests with the ”test_dummy_encryption” mount option. This option causes all new files to be automatically encrypted with a dummy key, without having to make any API calls. This tests the encrypted I/O paths more thoroughly. To do this with kvm-xfstests, use the ”encrypt” filesystem configuration:

```
kvm-xfstests -c ext4/encrypt,f2fs/encrypt -g auto
kvm-xfstests -c ext4/encrypt,f2fs/encrypt -g auto -m inlinecrypt
```

Because this runs many more tests than ”-g encrypt” does, it takes much longer to run; so also consider using gce-xfstests instead of kvm-xfstests:

```
gce-xfstests -c ext4/encrypt,f2fs/encrypt -g auto
gce-xfstests -c ext4/encrypt,f2fs/encrypt -g auto -m inlinecrypt
```

## 2.3 fs-verity: read-only file-based authenticity protection

### 2.3.1 Introduction

fs-verity (`fs/verity/`) is a support layer that filesystems can hook into to support transparent integrity and authenticity protection of read-only files. Currently, it is supported by the ext4, f2fs, and btrfs filesystems. Like fscrypt, not too much filesystem-specific code is needed to support fs-verity.

fs-verity is similar to dm-verity but works on files rather than block devices. On regular files on filesystems supporting fs-verity, userspace can execute an ioctl that causes the filesystem to build a Merkle tree for the file and persist it to a filesystem-specific location associated with the file.

After this, the file is made readonly, and all reads from the file are automatically verified against the file's Merkle tree. Reads of any corrupted data, including mmap reads, will fail.

Userspace can use another ioctl to retrieve the root hash (actually the "fs-verity file digest", which is a hash that includes the Merkle tree root hash) that fs-verity is enforcing for the file. This ioctl executes in constant time, regardless of the file size.

fs-verity is essentially a way to hash a file in constant time, subject to the caveat that reads which would violate the hash will fail at runtime.

### 2.3.2 Use cases

By itself, fs-verity only provides integrity protection, i.e. detection of accidental (non-malicious) corruption.

However, because fs-verity makes retrieving the file hash extremely efficient, it's primarily meant to be used as a tool to support authentication (detection of malicious modifications) or auditing (logging file hashes before use).

A standard file hash could be used instead of fs-verity. However, this is inefficient if the file is large and only a small portion may be accessed. This is often the case for Android application package (APK) files, for example. These typically contain many translations, classes, and other resources that are infrequently or even never accessed on a particular device. It would be slow and wasteful to read and hash the entire file before starting the application.

Unlike an ahead-of-time hash, fs-verity also re-verifies data each time it's paged in. This ensures that malicious disk firmware can't undetectably change the contents of the file at runtime.

fs-verity does not replace or obsolete dm-verity. dm-verity should still be used on read-only filesystems. fs-verity is for files that must live on a read-write filesystem because they are independently updated and potentially user-installed, so dm-verity cannot be used.

fs-verity does not mandate a particular scheme for authenticating its file hashes. (Similarly, dm-verity does not mandate a particular scheme for authenticating its block device root hashes.) Options for authenticating fs-verity file hashes include:

- Trusted userspace code. Often, the userspace code that accesses files can be trusted to authenticate them. Consider e.g. an application that wants to authenticate data files before using them, or an application loader that is part of the operating system (which is already authenticated in a different way, such as by being loaded from a read-only partition that uses dm-verity) and that wants to authenticate applications before loading them. In these cases, this trusted userspace code can authenticate a file's contents by retrieving its fs-verity digest using *FS_IOC_MEASURE_VERITY*, then verifying a signature of it using any userspace cryptographic library that supports digital signatures.
- Integrity Measurement Architecture (IMA). IMA supports fs-verity file digests as an alternative to its traditional full file digests. "IMA appraisal" enforces that files contain a valid, matching signature in their "security.ima" extended attribute, as controlled by the IMA policy. For more information, see the IMA documentation.
- Trusted userspace code in combination with *Built-in signature verification*. This approach should be used only with great care.

### 2.3.3 User API

#### FS_IOC_ENABLE_VERITY

The FS_IOC_ENABLE_VERITY ioctl enables fs-verity on a file. It takes in a pointer to a struct fsverity_enable_arg, defined as follows:

```
struct fsverity_enable_arg {
        __u32 version;
        __u32 hash_algorithm;
        __u32 block_size;
        __u32 salt_size;
        __u64 salt_ptr;
        __u32 sig_size;
        __u32 __reserved1;
        __u64 sig_ptr;
        __u64 __reserved2[11];
};
```

This structure contains the parameters of the Merkle tree to build for the file. It must be initialized as follows:

- `version` must be 1.
- `hash_algorithm` must be the identifier for the hash algorithm to use for the Merkle tree, such as FS_VERITY_HASH_ALG_SHA256. See `include/uapi/linux/fsverity.h` for the list of possible values.
- `block_size` is the Merkle tree block size, in bytes. In Linux v6.3 and later, this can be any power of 2 between (inclusively) 1024 and the minimum of the system page size and the filesystem block size. In earlier versions, the page size was the only allowed value.
- `salt_size` is the size of the salt in bytes, or 0 if no salt is provided. The salt is a value that is prepended to every hashed block; it can be used to personalize the hashing for a particular file or device. Currently the maximum salt size is 32 bytes.
- `salt_ptr` is the pointer to the salt, or NULL if no salt is provided.

- `sig_size` is the size of the builtin signature in bytes, or 0 if no builtin signature is provided. Currently the builtin signature is (somewhat arbitrarily) limited to 16128 bytes.
- `sig_ptr` is the pointer to the builtin signature, or NULL if no builtin signature is provided. A builtin signature is only needed if the *Built-in signature verification* feature is being used. It is not needed for IMA appraisal, and it is not needed if the file signature is being handled entirely in userspace.
- All reserved fields must be zeroed.

FS_IOC_ENABLE_VERITY causes the filesystem to build a Merkle tree for the file and persist it to a filesystem-specific location associated with the file, then mark the file as a verity file. This ioctl may take a long time to execute on large files, and it is interruptible by fatal signals.

FS_IOC_ENABLE_VERITY checks for write access to the inode. However, it must be executed on an O_RDONLY file descriptor and no processes can have the file open for writing. Attempts to open the file for writing while this ioctl is executing will fail with ETXTBSY. (This is necessary to guarantee that no writable file descriptors will exist after verity is enabled, and to guarantee that the file's contents are stable while the Merkle tree is being built over it.)

On success, FS_IOC_ENABLE_VERITY returns 0, and the file becomes a verity file. On failure (including the case of interruption by a fatal signal), no changes are made to the file.

FS_IOC_ENABLE_VERITY can fail with the following errors:

- `EACCES`: the process does not have write access to the file
- `EBADMSG`: the builtin signature is malformed
- `EBUSY`: this ioctl is already running on the file
- `EEXIST`: the file already has verity enabled
- `EFAULT`: the caller provided inaccessible memory
- `EFBIG`: the file is too large to enable verity on
- `EINTR`: the operation was interrupted by a fatal signal
- `EINVAL`: unsupported version, hash algorithm, or block size; or reserved bits are set; or the file descriptor refers to neither a regular file nor a directory.
- `EISDIR`: the file descriptor refers to a directory
- `EKEYREJECTED`: the builtin signature doesn't match the file
- `EMSGSIZE`: the salt or builtin signature is too long
- `ENOKEY`: the ".fs-verity" keyring doesn't contain the certificate needed to verify the builtin signature
- `ENOPKG`: fs-verity recognizes the hash algorithm, but it's not available in the kernel's crypto API as currently configured (e.g. for SHA-512, missing CONFIG_CRYPTO_SHA512).
- `ENOTTY`: this type of filesystem does not implement fs-verity
- `EOPNOTSUPP`: the kernel was not configured with fs-verity support; or the filesystem superblock has not had the 'verity' feature enabled on it; or the filesystem does not support fs-verity on this file. (See *Filesystem support*.)
- `EPERM`: the file is append-only; or, a builtin signature is required and one was not provided.
- `EROFS`: the filesystem is read-only

- ETXTBSY: someone has the file open for writing. This can be the caller's file descriptor, another open file descriptor, or the file reference held by a writable memory map.

### FS_IOC_MEASURE_VERITY

The FS_IOC_MEASURE_VERITY ioctl retrieves the digest of a verity file. The fs-verity file digest is a cryptographic digest that identifies the file contents that are being enforced on reads; it is computed via a Merkle tree and is different from a traditional full-file digest.

This ioctl takes in a pointer to a variable-length structure:

```
struct fsverity_digest {
        __u16 digest_algorithm;
        __u16 digest_size; /* input/output */
        __u8 digest[];
};
```

`digest_size` is an input/output field. On input, it must be initialized to the number of bytes allocated for the variable-length `digest` field.

On success, 0 is returned and the kernel fills in the structure as follows:

- `digest_algorithm` will be the hash algorithm used for the file digest. It will match `fsverity_enable_arg::hash_algorithm`.
- `digest_size` will be the size of the digest in bytes, e.g. 32 for SHA-256. (This can be redundant with `digest_algorithm`.)
- `digest` will be the actual bytes of the digest.

FS_IOC_MEASURE_VERITY is guaranteed to execute in constant time, regardless of the size of the file.

FS_IOC_MEASURE_VERITY can fail with the following errors:

- EFAULT: the caller provided inaccessible memory
- ENODATA: the file is not a verity file
- ENOTTY: this type of filesystem does not implement fs-verity
- EOPNOTSUPP: the kernel was not configured with fs-verity support, or the filesystem superblock has not had the 'verity' feature enabled on it. (See *Filesystem support*.)
- EOVERFLOW: the digest is longer than the specified `digest_size` bytes. Try providing a larger buffer.

### FS_IOC_READ_VERITY_METADATA

The FS_IOC_READ_VERITY_METADATA ioctl reads verity metadata from a verity file. This ioctl is available since Linux v5.12.

This ioctl allows writing a server program that takes a verity file and serves it to a client program, such that the client can do its own fs-verity compatible verification of the file. This only makes sense if the client doesn't trust the server and if the server needs to provide the storage for the client.

This is a fairly specialized use case, and most fs-verity users won't need this ioctl.

This ioctl takes in a pointer to the following structure:

```
#define FS_VERITY_METADATA_TYPE_MERKLE_TREE     1
#define FS_VERITY_METADATA_TYPE_DESCRIPTOR      2
#define FS_VERITY_METADATA_TYPE_SIGNATURE       3

struct fsverity_read_metadata_arg {
        __u64 metadata_type;
        __u64 offset;
        __u64 length;
        __u64 buf_ptr;
        __u64 __reserved;
};
```

`metadata_type` specifies the type of metadata to read:

- `FS_VERITY_METADATA_TYPE_MERKLE_TREE` reads the blocks of the Merkle tree. The blocks are returned in order from the root level to the leaf level. Within each level, the blocks are returned in the same order that their hashes are themselves hashed. See *Merkle tree* for more information.
- `FS_VERITY_METADATA_TYPE_DESCRIPTOR` reads the fs-verity descriptor. See *fs-verity descriptor*.
- `FS_VERITY_METADATA_TYPE_SIGNATURE` reads the builtin signature which was passed to FS_IOC_ENABLE_VERITY, if any. See *Built-in signature verification*.

The semantics are similar to those of `pread()`. `offset` specifies the offset in bytes into the metadata item to read from, and `length` specifies the maximum number of bytes to read from the metadata item. `buf_ptr` is the pointer to the buffer to read into, cast to a 64-bit integer. `__reserved` must be 0. On success, the number of bytes read is returned. 0 is returned at the end of the metadata item. The returned length may be less than `length`, for example if the ioctl is interrupted.

The metadata returned by FS_IOC_READ_VERITY_METADATA isn't guaranteed to be authenticated against the file digest that would be returned by *FS_IOC_MEASURE_VERITY*, as the metadata is expected to be used to implement fs-verity compatible verification anyway (though absent a malicious disk, the metadata will indeed match). E.g. to implement this ioctl, the filesystem is allowed to just read the Merkle tree blocks from disk without actually verifying the path to the root node.

FS_IOC_READ_VERITY_METADATA can fail with the following errors:

- `EFAULT`: the caller provided inaccessible memory

- `EINTR`: the ioctl was interrupted before any data was read
- `EINVAL`: reserved fields were set, or `offset + length` overflowed
- `ENODATA`: the file is not a verity file, or FS_VERITY_METADATA_TYPE_SIGNATURE was requested but the file doesn't have a builtin signature
- `ENOTTY`: this type of filesystem does not implement fs-verity, or this ioctl is not yet implemented on it
- `EOPNOTSUPP`: the kernel was not configured with fs-verity support, or the filesystem superblock has not had the 'verity' feature enabled on it. (See *Filesystem support*.)

#### FS_IOC_GETFLAGS

The existing ioctl FS_IOC_GETFLAGS (which isn't specific to fs-verity) can also be used to check whether a file has fs-verity enabled or not. To do so, check for FS_VERITY_FL (0x00100000) in the returned flags.

The verity flag is not settable via FS_IOC_SETFLAGS. You must use FS_IOC_ENABLE_VERITY instead, since parameters must be provided.

#### statx

Since Linux v5.5, the statx() system call sets STATX_ATTR_VERITY if the file has fs-verity enabled. This can perform better than FS_IOC_GETFLAGS and FS_IOC_MEASURE_VERITY because it doesn't require opening the file, and opening verity files can be expensive.

### 2.3.4 Accessing verity files

Applications can transparently access a verity file just like a non-verity one, with the following exceptions:

- Verity files are readonly. They cannot be opened for writing or truncate()d, even if the file mode bits allow it. Attempts to do one of these things will fail with EPERM. However, changes to metadata such as owner, mode, timestamps, and xattrs are still allowed, since these are not measured by fs-verity. Verity files can also still be renamed, deleted, and linked to.
- Direct I/O is not supported on verity files. Attempts to use direct I/O on such files will fall back to buffered I/O.
- DAX (Direct Access) is not supported on verity files, because this would circumvent the data verification.
- Reads of data that doesn't match the verity Merkle tree will fail with EIO (for read()) or SIGBUS (for mmap() reads).
- If the sysctl "fs.verity.require_signatures" is set to 1 and the file is not signed by a key in the ".fs-verity" keyring, then opening the file will fail. See *Built-in signature verification*.

Direct access to the Merkle tree is not supported. Therefore, if a verity file is copied, or is backed up and restored, then it will lose its "verity"-ness. fs-verity is primarily meant for files like executables that are managed by a package manager.

### 2.3.5 File digest computation

This section describes how fs-verity hashes the file contents using a Merkle tree to produce the digest which cryptographically identifies the file contents. This algorithm is the same for all filesystems that support fs-verity.

Userspace only needs to be aware of this algorithm if it needs to compute fs-verity file digests itself, e.g. in order to sign files.

#### Merkle tree

The file contents is divided into blocks, where the block size is configurable but is usually 4096 bytes. The end of the last block is zero-padded if needed. Each block is then hashed, producing the first level of hashes. Then, the hashes in this first level are grouped into 'blocksize'-byte blocks (zero-padding the ends as needed) and these blocks are hashed, producing the second level of hashes. This proceeds up the tree until only a single block remains. The hash of this block is the ”Merkle tree root hash”.

If the file fits in one block and is nonempty, then the ”Merkle tree root hash” is simply the hash of the single data block. If the file is empty, then the ”Merkle tree root hash” is all zeroes.

The ”blocks” here are not necessarily the same as ”filesystem blocks”.

If a salt was specified, then it's zero-padded to the closest multiple of the input size of the hash algorithm's compression function, e.g. 64 bytes for SHA-256 or 128 bytes for SHA-512. The padded salt is prepended to every data or Merkle tree block that is hashed.

The purpose of the block padding is to cause every hash to be taken over the same amount of data, which simplifies the implementation and keeps open more possibilities for hardware acceleration. The purpose of the salt padding is to make the salting ”free” when the salted hash state is precomputed, then imported for each hash.

Example: in the recommended configuration of SHA-256 and 4K blocks, 128 hash values fit in each block. Thus, each level of the Merkle tree is approximately 128 times smaller than the previous, and for large files the Merkle tree's size converges to approximately 1/127 of the original file size. However, for small files, the padding is significant, making the space overhead proportionally more.

#### fs-verity descriptor

By itself, the Merkle tree root hash is ambiguous. For example, it can't a distinguish a large file from a small second file whose data is exactly the top-level hash block of the first file. Ambiguities also arise from the convention of padding to the next block boundary.

To solve this problem, the fs-verity file digest is actually computed as a hash of the following structure, which contains the Merkle tree root hash as well as other fields such as the file size:

```
struct fsverity_descriptor {
        __u8 version;           /* must be 1 */
        __u8 hash_algorithm;    /* Merkle tree hash algorithm */
        __u8 log_blocksize;     /* log2 of size of data and tree blocks */
        __u8 salt_size;         /* size of salt in bytes; 0 if none */
        __le32 __reserved_0x04; /* must be 0 */
```

```
        __le64 data_size;       /* size of file the Merkle tree is built over␣
↪*/
        __u8 root_hash[64];     /* Merkle tree root hash */
        __u8 salt[32];          /* salt prepended to each hashed block */
        __u8 __reserved[144];   /* must be 0's */
};
```

### 2.3.6 Built-in signature verification

CONFIG_FS_VERITY_BUILTIN_SIGNATURES=y adds supports for in-kernel verification of fs-verity builtin signatures.

**IMPORTANT**! Please take great care before using this feature. It is not the only way to do signatures with fs-verity, and the alternatives (such as userspace signature verification, and IMA appraisal) can be much better. It's also easy to fall into a trap of thinking this feature solves more problems than it actually does.

Enabling this option adds the following:

1. At boot time, the kernel creates a keyring named ".fs-verity". The root user can add trusted X.509 certificates to this keyring using the add_key() system call.
2. *FS_IOC_ENABLE_VERITY* accepts a pointer to a PKCS#7 formatted detached signature in DER format of the file's fs-verity digest. On success, the ioctl persists the signature alongside the Merkle tree. Then, any time the file is opened, the kernel verifies the file's actual digest against this signature, using the certificates in the ".fs-verity" keyring.
3. A new sysctl "fs.verity.require_signatures" is made available. When set to 1, the kernel requires that all verity files have a correctly signed digest as described in (2).

The data that the signature as described in (2) must be a signature of is the fs-verity file digest in the following format:

```
struct fsverity_formatted_digest {
        char magic[8];                  /* must be "FSVerity" */
        __le16 digest_algorithm;
        __le16 digest_size;
        __u8 digest[];
};
```

That's it. It should be emphasized again that fs-verity builtin signatures are not the only way to do signatures with fs-verity. See *Use cases* for an overview of ways in which fs-verity can be used. fs-verity builtin signatures have some major limitations that should be carefully considered before using them:

- Builtin signature verification does *not* make the kernel enforce that any files actually have fs-verity enabled. Thus, it is not a complete authentication policy. Currently, if it is used, the only way to complete the authentication policy is for trusted userspace code to explicitly check whether files have fs-verity enabled with a signature before they are accessed. (With fs.verity.require_signatures=1, just checking whether fs-verity is enabled suffices.) But, in this case the trusted userspace code could just store the signature alongside the file and verify it itself using a cryptographic library, instead of using this feature.

- A file's builtin signature can only be set at the same time that fs-verity is being enabled on the file. Changing or deleting the builtin signature later requires re-creating the file.
- Builtin signature verification uses the same set of public keys for all fs-verity enabled files on the system. Different keys cannot be trusted for different files; each key is all or nothing.
- The sysctl fs.verity.require_signatures applies system-wide. Setting it to 1 only works when all users of fs-verity on the system agree that it should be set to 1. This limitation can prevent fs-verity from being used in cases where it would be helpful.
- Builtin signature verification can only use signature algorithms that are supported by the kernel. For example, the kernel does not yet support Ed25519, even though this is often the signature algorithm that is recommended for new cryptographic designs.
- fs-verity builtin signatures are in PKCS#7 format, and the public keys are in X.509 format. These formats are commonly used, including by some other kernel features (which is why the fs-verity builtin signatures use them), and are very feature rich. Unfortunately, history has shown that code that parses and handles these formats (which are from the 1990s and are based on ASN.1) often has vulnerabilities as a result of their complexity. This complexity is not inherent to the cryptography itself.

  fs-verity users who do not need advanced features of X.509 and PKCS#7 should strongly consider using simpler formats, such as plain Ed25519 keys and signatures, and verifying signatures in userspace.

  fs-verity users who choose to use X.509 and PKCS#7 anyway should still consider that verifying those signatures in userspace is more flexible (for other reasons mentioned earlier in this document) and eliminates the need to enable CONFIG_FS_VERITY_BUILTIN_SIGNATURES and its associated increase in kernel attack surface. In some cases it can even be necessary, since advanced X.509 and PKCS#7 features do not always work as intended with the kernel. For example, the kernel does not check X.509 certificate validity times.

  Note: IMA appraisal, which supports fs-verity, does not use PKCS#7 for its signatures, so it partially avoids the issues discussed here. IMA appraisal does use X.509.

### 2.3.7 Filesystem support

fs-verity is supported by several filesystems, described below. The CONFIG_FS_VERITY kconfig option must be enabled to use fs-verity on any of these filesystems.

`include/linux/fsverity.h` declares the interface between the `fs/verity/` support layer and filesystems. Briefly, filesystems must provide an `fsverity_operations` structure that provides methods to read and write the verity metadata to a filesystem-specific location, including the Merkle tree blocks and `fsverity_descriptor`. Filesystems must also call functions in `fs/verity/` at certain times, such as when a file is opened or when pages have been read into the pagecache. (See *Verifying data*.)

### ext4

ext4 supports fs-verity since Linux v5.4 and e2fsprogs v1.45.2.

To create verity files on an ext4 filesystem, the filesystem must have been formatted with `-O verity` or had `tune2fs -O verity` run on it. "verity" is an RO_COMPAT filesystem feature, so once set, old kernels will only be able to mount the filesystem readonly, and old versions of e2fsck will be unable to check the filesystem.

Originally, an ext4 filesystem with the "verity" feature could only be mounted when its block size was equal to the system page size (typically 4096 bytes). In Linux v6.3, this limitation was removed.

ext4 sets the EXT4_VERITY_FL on-disk inode flag on verity files. It can only be set by *FS_IOC_ENABLE_VERITY*, and it cannot be cleared.

ext4 also supports encryption, which can be used simultaneously with fs-verity. In this case, the plaintext data is verified rather than the ciphertext. This is necessary in order to make the fs-verity file digest meaningful, since every file is encrypted differently.

ext4 stores the verity metadata (Merkle tree and fsverity_descriptor) past the end of the file, starting at the first 64K boundary beyond i_size. This approach works because (a) verity files are readonly, and (b) pages fully beyond i_size aren't visible to userspace but can be read/written internally by ext4 with only some relatively small changes to ext4. This approach avoids having to depend on the EA_INODE feature and on rearchitecturing ext4's xattr support to support paging multi-gigabyte xattrs into memory, and to support encrypting xattrs. Note that the verity metadata *must* be encrypted when the file is, since it contains hashes of the plaintext data.

ext4 only allows verity on extent-based files.

### f2fs

f2fs supports fs-verity since Linux v5.4 and f2fs-tools v1.11.0.

To create verity files on an f2fs filesystem, the filesystem must have been formatted with `-O verity`.

f2fs sets the FADVISE_VERITY_BIT on-disk inode flag on verity files. It can only be set by *FS_IOC_ENABLE_VERITY*, and it cannot be cleared.

Like ext4, f2fs stores the verity metadata (Merkle tree and fsverity_descriptor) past the end of the file, starting at the first 64K boundary beyond i_size. See explanation for ext4 above. Moreover, f2fs supports at most 4096 bytes of xattr entries per inode which usually wouldn't be enough for even a single Merkle tree block.

f2fs doesn't support enabling verity on files that currently have atomic or volatile writes pending.

#### btrfs

btrfs supports fs-verity since Linux v5.15. Verity-enabled inodes are marked with a RO_COMPAT inode flag, and the verity metadata is stored in separate btree items.

### 2.3.8 Implementation details

#### Verifying data

fs-verity ensures that all reads of a verity file's data are verified, regardless of which syscall is used to do the read (e.g. mmap(), read(), pread()) and regardless of whether it's the first read or a later read (unless the later read can return cached data that was already verified). Below, we describe how filesystems implement this.

#### Pagecache

For filesystems using Linux's pagecache, the `->read_folio()` and `->readahead()` methods must be modified to verify folios before they are marked Uptodate. Merely hooking `->read_iter()` would be insufficient, since `->read_iter()` is not used for memory maps.

Therefore, fs/verity/ provides the function fsverity_verify_blocks() which verifies data that has been read into the pagecache of a verity inode. The containing folio must still be locked and not Uptodate, so it's not yet readable by userspace. As needed to do the verification, fsverity_verify_blocks() will call back into the filesystem to read hash blocks via fsverity_operations::read_merkle_tree_page().

fsverity_verify_blocks() returns false if verification failed; in this case, the filesystem must not set the folio Uptodate. Following this, as per the usual Linux pagecache behavior, attempts by userspace to read() from the part of the file containing the folio will fail with EIO, and accesses to the folio within a memory map will raise SIGBUS.

In principle, verifying a data block requires verifying the entire path in the Merkle tree from the data block to the root hash. However, for efficiency the filesystem may cache the hash blocks. Therefore, fsverity_verify_blocks() only ascends the tree reading hash blocks until an already-verified hash block is seen. It then verifies the path to that block.

This optimization, which is also used by dm-verity, results in excellent sequential read performance. This is because usually (e.g. 127 in 128 times for 4K blocks and SHA-256) the hash block from the bottom level of the tree will already be cached and checked from reading a previous data block. However, random reads perform worse.

#### Block device based filesystems

Block device based filesystems (e.g. ext4 and f2fs) in Linux also use the pagecache, so the above subsection applies too. However, they also usually read many data blocks from a file at once, grouped into a structure called a "bio". To make it easier for these types of filesystems to support fs-verity, fs/verity/ also provides a function fsverity_verify_bio() which verifies all data blocks in a bio.

ext4 and f2fs also support encryption. If a verity file is also encrypted, the data must be decrypted before being verified. To support this, these filesystems allocate a "post-read context" for each bio and store it in `->bi_private`:

```
struct bio_post_read_ctx {
       struct bio *bio;
       struct work_struct work;
       unsigned int cur_step;
       unsigned int enabled_steps;
};
```

`enabled_steps` is a bitmask that specifies whether decryption, verity, or both is enabled. After the bio completes, for each needed postprocessing step the filesystem enqueues the bio_post_read_ctx on a workqueue, and then the workqueue work does the decryption or verification. Finally, folios where no decryption or verity error occurred are marked Uptodate, and the folios are unlocked.

On many filesystems, files can contain holes. Normally, `->readahead()` simply zeroes hole blocks and considers the corresponding data to be up-to-date; no bios are issued. To prevent this case from bypassing fs-verity, filesystems use fsverity_verify_blocks() to verify hole blocks.

Filesystems also disable direct I/O on verity files, since otherwise direct I/O would bypass fs-verity.

### 2.3.9 Userspace utility

This document focuses on the kernel, but a userspace utility for fs-verity can be found at:

https://git.kernel.org/pub/scm/fs/fsverity/fsverity-utils.git

See the README.md file in the fsverity-utils source tree for details, including examples of setting up fs-verity protected files.

### 2.3.10 Tests

To test fs-verity, use xfstests. For example, using kvm-xfstests:

```
kvm-xfstests -c ext4,f2fs,btrfs -g verity
```

### 2.3.11 FAQ

This section answers frequently asked questions about fs-verity that weren't already directly answered in other parts of this document.

**Q**

Why isn't fs-verity part of IMA?

**A**

fs-verity and IMA (Integrity Measurement Architecture) have different focuses. fs-verity is a filesystem-level mechanism for hashing individual files using a Merkle tree. In contrast, IMA specifies a system-wide policy that specifies which

files are hashed and what to do with those hashes, such as log them, authenticate them, or add them to a measurement list.

IMA supports the fs-verity hashing mechanism as an alternative to full file hashes, for those who want the performance and security benefits of the Merkle tree based hash. However, it doesn't make sense to force all uses of fs-verity to be through IMA. fs-verity already meets many users' needs even as a standalone filesystem feature, and it's testable like other filesystem features e.g. with xfstests.

**Q**

Isn't fs-verity useless because the attacker can just modify the hashes in the Merkle tree, which is stored on-disk?

**A**

To verify the authenticity of an fs-verity file you must verify the authenticity of the "fs-verity file digest", which incorporates the root hash of the Merkle tree. See *Use cases*.

**Q**

Isn't fs-verity useless because the attacker can just replace a verity file with a non-verity one?

**A**

See *Use cases*. In the initial use case, it's really trusted userspace code that authenticates the files; fs-verity is just a tool to do this job efficiently and securely. The trusted userspace code will consider non-verity files to be inauthentic.

**Q**

Why does the Merkle tree need to be stored on-disk? Couldn't you store just the root hash?

**A**

If the Merkle tree wasn't stored on-disk, then you'd have to compute the entire tree when the file is first accessed, even if just one byte is being read. This is a fundamental consequence of how Merkle tree hashing works. To verify a leaf node, you need to verify the whole path to the root hash, including the root node (the thing which the root hash is a hash of). But if the root node isn't stored on-disk, you have to compute it by hashing its children, and so on until you've actually hashed the entire file.

That defeats most of the point of doing a Merkle tree-based hash, since if you have to hash the whole file ahead of time anyway, then you could simply do sha256(file) instead. That would be much simpler, and a bit faster too.

It's true that an in-memory Merkle tree could still provide the advantage of verification on every read rather than just on the first read. However, it would be inefficient because every time a hash page gets evicted (you can't pin the entire Merkle tree into memory, since it may be very large), in order to restore it you again need to hash everything below it in the tree. This again defeats most of the point of doing a Merkle tree-based hash, since a single block read could trigger re-hashing gigabytes of data.

**Q**

But couldn't you store just the leaf nodes and compute the rest?

**A**

See previous answer; this really just moves up one level, since one could alternatively interpret the data blocks as being the leaf nodes of the Merkle tree. It's true that the tree can be computed much faster if the leaf level is stored rather than just the data, but that's only because each level is less than 1% the size of the level below (assuming the recommended settings of SHA-256 and 4K blocks). For the exact same reason, by storing "just the leaf nodes" you'd already be storing over 99% of the tree, so you might as well simply store the whole tree.

**Q**

Can the Merkle tree be built ahead of time, e.g. distributed as part of a package that is installed to many computers?

**A**

This isn't currently supported. It was part of the original design, but was removed to simplify the kernel UAPI and because it wasn't a critical use case. Files are usually installed once and used many times, and cryptographic hashing is somewhat fast on most modern processors.

**Q**

Why doesn't fs-verity support writes?

**A**

Write support would be very difficult and would require a completely different design, so it's well outside the scope of fs-verity. Write support would require:

- A way to maintain consistency between the data and hashes, including all levels of hashes, since corruption after a crash (especially of potentially the entire file!) is unacceptable. The main options for solving this are data journalling, copy-on-write, and log-structured volume. But it's very hard to retrofit existing filesystems with new consistency mechanisms. Data journalling is available on ext4, but is very slow.
- Rebuilding the Merkle tree after every write, which would be extremely inefficient. Alternatively, a different authenticated dictionary structure such as an "authenticated skiplist" could be used. However, this would be far more complex.

Compare it to dm-verity vs. dm-integrity. dm-verity is very simple: the kernel just verifies read-only data against a read-only Merkle tree. In contrast, dm-integrity supports writes but is slow, is much more complex, and doesn't actually support full-device authentication since it authenticates each sector independently, i.e. there is no "root hash". It doesn't really make sense for the same device-mapper target to support these two very different cases; the same applies to fs-verity.

**Q**

Since verity files are immutable, why isn't the immutable bit set?

**A**

The existing "immutable" bit (FS_IMMUTABLE_FL) already has a specific set of semantics which not only make the file contents read-only, but also prevent the file from being deleted, renamed, linked to, or having its owner or mode changed. These extra properties are unwanted for fs-verity, so reusing the immutable bit isn't appropriate.

**Q**

Why does the API use ioctls instead of setxattr() and getxattr()?

**A**

Abusing the xattr interface for basically arbitrary syscalls is heavily frowned upon by most of the Linux filesystem developers. An xattr should really just be an xattr on-disk, not an API to e.g. magically trigger construction of a Merkle tree.

**Q**

Does fs-verity support remote filesystems?

**A**

So far all filesystems that have implemented fs-verity support are local filesystems, but in principle any filesystem that can store per-file verity metadata can support fs-verity, regardless of whether it's local or remote. Some filesystems may have fewer options of where to store the verity metadata; one possibility is to store it past the end of the file and "hide" it from userspace by manipulating i_size. The data verification functions provided by `fs/verity/` also assume that the filesystem uses the Linux pagecache, but both local and remote filesystems normally do so.

**Q**

Why is anything filesystem-specific at all? Shouldn't fs-verity be implemented entirely at the VFS level?

**A**

There are many reasons why this is not possible or would be very difficult, including the following:

- To prevent bypassing verification, folios must not be marked Uptodate until they've been verified. Currently, each filesystem is responsible for marking folios Uptodate via `->readahead()`. Therefore, currently it's not possible for the VFS to do the verification on its own. Changing this would require significant changes to the VFS and all filesystems.
- It would require defining a filesystem-independent way to store the verity metadata. Extended attributes don't work for this because (a) the Merkle tree may be gigabytes, but many filesystems assume that all xattrs fit into a single 4K filesystem block, and (b) ext4 and f2fs encryption doesn't encrypt xattrs, yet the Merkle tree *must* be encrypted when the file contents are, because it stores hashes of the plaintext file contents.

  So the verity metadata would have to be stored in an actual file. Using a separate file would be very ugly, since the metadata is fundamentally part of the file to be protected, and it could cause problems where users could delete the real file but not the metadata file or vice versa. On the other hand, having it be in the same file would break applications unless filesystems' notion of i_size were divorced from the VFS's, which would be complex and require changes to all filesystems.
- It's desirable that FS_IOC_ENABLE_VERITY uses the filesystem's transaction mechanism so that either the file ends up with verity enabled, or no changes were made. Allowing intermediate states to occur after a crash may cause problems.

## 2.4 Network Filesystem Helper Library

### 2.4.1 Overview

The network filesystem helper library is a set of functions designed to aid a network filesystem in implementing VM/VFS operations. For the moment, that just includes turning various VM buffered read operations into requests to read from the server. The helper library, however, can also interpose other services, such as local caching or local data encryption.

Note that the library module doesn't link against local caching directly, so access must be provided by the netfs.

### 2.4.2 Per-Inode Context

The network filesystem helper library needs a place to store a bit of state for its use on each netfs inode it is helping to manage. To this end, a context structure is defined:

```
struct netfs_inode {
        struct inode inode;
        const struct netfs_request_ops *ops;
        struct fscache_cookie *cache;
};
```

A network filesystem that wants to use netfs lib must place one of these in its inode wrapper struct instead of the VFS `struct inode`. This can be done in a way similar to the following:

```
struct my_inode {
        struct netfs_inode netfs; /* Netfslib context and vfs inode */
        ...
};
```

This allows netfslib to find its state by using `container_of()` from the inode pointer, thereby allowing the netfslib helper functions to be pointed to directly by the VFS/VM operation tables.

The structure contains the following fields:

- `inode`

  The VFS inode structure.

- `ops`

  The set of operations provided by the network filesystem to netfslib.

- `cache`

  Local caching cookie, or NULL if no caching is enabled. This field does not exist if fscache is disabled.

#### Inode Context Helper Functions

To help deal with the per-inode context, a number helper functions are provided. Firstly, a function to perform basic initialisation on a context and set the operations table pointer:

```
void netfs_inode_init(struct netfs_inode *ctx,
                      const struct netfs_request_ops *ops);
```

then a function to cast from the VFS inode structure to the netfs context:

```
struct netfs_inode *netfs_node(struct inode *inode);
```

and finally, a function to get the cache cookie pointer from the context attached to an inode (or NULL if fscache is disabled):

```
struct fscache_cookie *netfs_i_cookie(struct netfs_inode *ctx);
```

### 2.4.3 Buffered Read Helpers

The library provides a set of read helpers that handle the ->read_folio(), ->readahead() and much of the ->write_begin() VM operations and translate them into a common call framework.

The following services are provided:

- Handle folios that span multiple pages.
- Insulate the netfs from VM interface changes.
- Allow the netfs to arbitrarily split reads up into pieces, even ones that don't match folio sizes or folio alignments and that may cross folios.
- Allow the netfs to expand a readahead request in both directions to meet its needs.
- Allow the netfs to partially fulfil a read, which will then be resubmitted.
- Handle local caching, allowing cached data and server-read data to be interleaved for a single request.
- Handle clearing of bufferage that aren't on the server.
- Handle retrying of reads that failed, switching reads from the cache to the server as necessary.
- In the future, this is a place that other services can be performed, such as local encryption of data to be stored remotely or in the cache.

From the network filesystem, the helpers require a table of operations. This includes a mandatory method to issue a read operation along with a number of optional methods.

### Read Helper Functions

Three read helpers are provided:

```
void netfs_readahead(struct readahead_control *ractl);
int netfs_read_folio(struct file *file,
                     struct folio *folio);
int netfs_write_begin(struct netfs_inode *ctx,
                      struct file *file,
                      struct address_space *mapping,
                      loff_t pos,
                      unsigned int len,
                      struct folio **_folio,
                      void **_fsdata);
```

Each corresponds to a VM address space operation. These operations use the state in the per-inode context.

For ->readahead() and ->read_folio(), the network filesystem just point directly at the corresponding read helper; whereas for ->write_begin(), it may be a little more complicated as the network filesystem might want to flush conflicting writes or track dirty data and needs to put the acquired folio if an error occurs after calling the helper.

The helpers manage the read request, calling back into the network filesystem through the supplied table of operations. Waits will be performed as necessary before returning for helpers that are meant to be synchronous.

If an error occurs, the ->free_request() will be called to clean up the netfs_io_request struct allocated. If some parts of the request are in progress when an error occurs, the request will get partially completed if sufficient data is read.

Additionally, there is:

```
* void netfs_subreq_terminated(struct netfs_io_subrequest *subreq,
                               ssize_t transferred_or_error,
                               bool was_async);
```

which should be called to complete a read subrequest. This is given the number of bytes transferred or a negative error code, plus a flag indicating whether the operation was asynchronous (ie. whether the follow-on processing can be done in the current context, given this may involve sleeping).

### Read Helper Structures

The read helpers make use of a couple of structures to maintain the state of the read. The first is a structure that manages a read request as a whole:

```
struct netfs_io_request {
        struct inode            *inode;
        struct address_space    *mapping;
        struct netfs_cache_resources cache_resources;
        void                    *netfs_priv;
        loff_t                  start;
```

```
        size_t                  len;
        loff_t                  i_size;
        const struct netfs_request_ops *netfs_ops;
        unsigned int            debug_id;
        ...
};
```

The above fields are the ones the netfs can use. They are:

- `inode`
- `mapping`

  The inode and the address space of the file being read from. The mapping may or may not point to inode->i_data.

- `cache_resources`

  Resources for the local cache to use, if present.

- `netfs_priv`

  The network filesystem's private data. The value for this can be passed in to the helper functions or set during the request.

- `start`
- `len`

  The file position of the start of the read request and the length. These may be altered by the ->expand_readahead() op.

- `i_size`

  The size of the file at the start of the request.

- `netfs_ops`

  A pointer to the operation table. The value for this is passed into the helper functions.

- `debug_id`

  A number allocated to this operation that can be displayed in trace lines for reference.

The second structure is used to manage individual slices of the overall read request:

```
struct netfs_io_subrequest {
        struct netfs_io_request *rreq;
        loff_t                  start;
        size_t                  len;
        size_t                  transferred;
        unsigned long           flags;
        unsigned short          debug_index;
        ...
};
```

Each subrequest is expected to access a single source, though the helpers will handle falling back from one source type to another. The members are:

- `rreq`

  A pointer to the read request.

- `start`
- `len`

  The file position of the start of this slice of the read request and the length.

- `transferred`

  The amount of data transferred so far of the length of this slice. The network filesystem or cache should start the operation this far into the slice. If a short read occurs, the helpers will call again, having updated this to reflect the amount read so far.

- `flags`

  Flags pertaining to the read. There are two of interest to the filesystem or cache:

  - `NETFS_SREQ_CLEAR_TAIL`

    This can be set to indicate that the remainder of the slice, from transferred to len, should be cleared.

  - `NETFS_SREQ_SEEK_DATA_READ`

    This is a hint to the cache that it might want to try skipping ahead to the next data (ie. using SEEK_DATA).

- `debug_index`

  A number allocated to this slice that can be displayed in trace lines for reference.

### Read Helper Operations

The network filesystem must provide the read helpers with a table of operations through which it can issue requests and negotiate:

```
struct netfs_request_ops {
        void (*init_request)(struct netfs_io_request *rreq, struct file *file);
        void (*free_request)(struct netfs_io_request *rreq);
        void (*expand_readahead)(struct netfs_io_request *rreq);
        bool (*clamp_length)(struct netfs_io_subrequest *subreq);
        void (*issue_read)(struct netfs_io_subrequest *subreq);
        bool (*is_still_valid)(struct netfs_io_request *rreq);
        int (*check_write_begin)(struct file *file, loff_t pos, unsigned len,
                                 struct folio **foliop, void **_fsdata);
        void (*done)(struct netfs_io_request *rreq);
};
```

The operations are as follows:

- `init_request()`

  [Optional] This is called to initialise the request structure. It is given the file for reference.

- `free_request()`

[Optional] This is called as the request is being deallocated so that the filesystem can clean up any state it has attached there.

- `expand_readahead()`

  [Optional] This is called to allow the filesystem to expand the size of a readahead read request. The filesystem gets to expand the request in both directions, though it's not permitted to reduce it as the numbers may represent an allocation already made. If local caching is enabled, it gets to expand the request first.

  Expansion is communicated by changing ->start and ->len in the request structure. Note that if any change is made, ->len must be increased by at least as much as ->start is reduced.

- `clamp_length()`

  [Optional] This is called to allow the filesystem to reduce the size of a subrequest. The filesystem can use this, for example, to chop up a request that has to be split across multiple servers or to put multiple reads in flight.

  This should return 0 on success and an error code on error.

- `issue_read()`

  [Required] The helpers use this to dispatch a subrequest to the server for reading. In the subrequest, ->start, ->len and ->transferred indicate what data should be read from the server.

  There is no return value; the *netfs_subreq_terminated()* function should be called to indicate whether or not the operation succeeded and how much data it transferred. The filesystem also should not deal with setting folios uptodate, unlocking them or dropping their refs - the helpers need to deal with this as they have to coordinate with copying to the local cache.

  Note that the helpers have the folios locked, but not pinned. It is possible to use the ITER_XARRAY iov iterator to refer to the range of the inode that is being operated upon without the need to allocate large bvec tables.

- `is_still_valid()`

  [Optional] This is called to find out if the data just read from the local cache is still valid. It should return true if it is still valid and false if not. If it's not still valid, it will be reread from the server.

- `check_write_begin()`

  [Optional] This is called from the *netfs_write_begin()* helper once it has allocated/grabbed the folio to be modified to allow the filesystem to flush conflicting state before allowing it to be modified.

  It may unlock and discard the folio it was given and set the caller's folio pointer to NULL. It should return 0 if everything is now fine (`*foliop` left set) or the op should be retried (`*foliop` cleared) and any other error code to abort the operation.

- `done`

  [Optional] This is called after the folios in the request have all been unlocked (and marked uptodate if applicable).

### Read Helper Procedure

The read helpers work by the following general procedure:

- Set up the request.
- For readahead, allow the local cache and then the network filesystem to propose expansions to the read request. This is then proposed to the VM. If the VM cannot fully perform the expansion, a partially expanded read will be performed, though this may not get written to the cache in its entirety.
- Loop around slicing chunks off of the request to form subrequests:
  - If a local cache is present, it gets to do the slicing, otherwise the helpers just try to generate maximal slices.
  - The network filesystem gets to clamp the size of each slice if it is to be the source. This allows rsize and chunking to be implemented.
  - The helpers issue a read from the cache or a read from the server or just clears the slice as appropriate.
  - The next slice begins at the end of the last one.
  - As slices finish being read, they terminate.
- When all the subrequests have terminated, the subrequests are assessed and any that are short or have failed are reissued:
  - Failed cache requests are issued against the server instead.
  - Failed server requests just fail.
  - Short reads against either source will be reissued against that source provided they have transferred some more data:
    - The cache may need to skip holes that it can't do DIO from.
    - If NETFS_SREQ_CLEAR_TAIL was set, a short read will be cleared to the end of the slice instead of reissuing.
- Once the data is read, the folios that have been fully read/cleared:
  - Will be marked uptodate.
  - If a cache is present, will be marked with PG_fscache.
  - Unlocked
- Any folios that need writing to the cache will then have DIO writes issued.
- Synchronous operations will wait for reading to be complete.
- Writes to the cache will proceed asynchronously and the folios will have the PG_fscache mark removed when that completes.
- The request structures will be cleaned up when everything has completed.

### Read Helper Cache API

When implementing a local cache to be used by the read helpers, two things are required: some way for the network filesystem to initialise the caching for a read request and a table of operations for the helpers to call.

To begin a cache operation on an fscache object, the following function is called:

```
int fscache_begin_read_operation(struct netfs_io_request *rreq,
                                 struct fscache_cookie *cookie);
```

passing in the request pointer and the cookie corresponding to the file. This fills in the cache resources mentioned below.

The netfs_io_request object contains a place for the cache to hang its state:

```
struct netfs_cache_resources {
        const struct netfs_cache_ops    *ops;
        void                            *cache_priv;
        void                            *cache_priv2;
};
```

This contains an operations table pointer and two private pointers. The operation table looks like the following:

```
struct netfs_cache_ops {
        void (*end_operation)(struct netfs_cache_resources *cres);

        void (*expand_readahead)(struct netfs_cache_resources *cres,
                                 loff_t *_start, size_t *_len, loff_t i_size);

        enum netfs_io_source (*prepare_read)(struct netfs_io_subrequest␣
↪*subreq,
                                             loff_t i_size);

        int (*read)(struct netfs_cache_resources *cres,
                    loff_t start_pos,
                    struct iov_iter *iter,
                    bool seek_data,
                    netfs_io_terminated_t term_func,
                    void *term_func_priv);

        int (*prepare_write)(struct netfs_cache_resources *cres,
                             loff_t *_start, size_t *_len, loff_t i_size,
                             bool no_space_allocated_yet);

        int (*write)(struct netfs_cache_resources *cres,
                     loff_t start_pos,
                     struct iov_iter *iter,
                     netfs_io_terminated_t term_func,
                     void *term_func_priv);

        int (*query_occupancy)(struct netfs_cache_resources *cres,
```

```
                                loff_t start, size_t len, size_t granularity,
                                loff_t *_data_start, size_t *_data_len);
};
```

With a termination handler function pointer:

```
typedef void (*netfs_io_terminated_t)(void *priv,
                                      ssize_t transferred_or_error,
                                      bool was_async);
```

The methods defined in the table are:

- `end_operation()`

  [Required] Called to clean up the resources at the end of the read request.

- `expand_readahead()`

  [Optional] Called at the beginning of a *`netfs_readahead()`* operation to allow the cache to expand a request in either direction. This allows the cache to size the request appropriately for the cache granularity.

  The function is passed poiners to the start and length in its parameters, plus the size of the file for reference, and adjusts the start and length appropriately. It should return one of:

  - `NETFS_FILL_WITH_ZEROES`
  - `NETFS_DOWNLOAD_FROM_SERVER`
  - `NETFS_READ_FROM_CACHE`
  - `NETFS_INVALID_READ`

  to indicate whether the slice should just be cleared or whether it should be downloaded from the server or read from the cache - or whether slicing should be given up at the current point.

- `prepare_read()`

  [Required] Called to configure the next slice of a request. ->start and ->len in the subrequest indicate where and how big the next slice can be; the cache gets to reduce the length to match its granularity requirements.

- `read()`

  [Required] Called to read from the cache. The start file offset is given along with an iterator to read to, which gives the length also. It can be given a hint requesting that it seek forward from that start position for data.

  Also provided is a pointer to a termination handler function and private data to pass to that function. The termination function should be called with the number of bytes transferred or an error code, plus a flag indicating whether the termination is definitely happening in the caller's context.

- `prepare_write()`

  [Required] Called to prepare a write to the cache to take place. This involves checking to see whether the cache has sufficient space to honour the write. `*_start` and `*_len`

indicate the region to be written; the region can be shrunk or it can be expanded to a page boundary either way as necessary to align for direct I/O. i_size holds the size of the object and is provided for reference. no_space_allocated_yet is set to true if the caller is certain that no data has been written to that region - for example if it tried to do a read from there already.

- `write()`

  [Required] Called to write to the cache. The start file offset is given along with an iterator to write from, which gives the length also.

  Also provided is a pointer to a termination handler function and private data to pass to that function. The termination function should be called with the number of bytes transferred or an error code, plus a flag indicating whether the termination is definitely happening in the caller's context.

- `query_occupancy()`

  [Required] Called to find out where the next piece of data is within a particular region of the cache. The start and length of the region to be queried are passed in, along with the granularity to which the answer needs to be aligned. The function passes back the start and length of the data, if any, available within that region. Note that there may be a hole at the front.

  It returns 0 if some data was found, -ENODATA if there was no usable data within the region or -ENOBUFS if there is no caching on this file.

Note that these methods are passed a pointer to the cache resource structure, not the read request structure as they could be used in other situations where there isn't a read request structure as well, such as writing dirty data to the cache.

### 2.4.4 API Function Reference

void **folio_start_fscache**(struct *folio* *folio)

Start an fscache write on a folio.

**Parameters**

**struct folio *folio**
The folio.

**Description**

Call this function before writing a folio to a local cache. Starting a second write before the first one finishes is not allowed.

void **folio_end_fscache**(struct *folio* *folio)

End an fscache write on a folio.

**Parameters**

**struct folio *folio**
The folio.

**Description**

Call this function after the folio has been written to the local cache. This will wake any sleepers waiting on this folio.

void **folio_wait_fscache**(struct *folio* *folio)

Wait for an fscache write on this folio to end.

**Parameters**

**struct folio *folio**

The folio.

**Description**

If this folio is currently being written to a local cache, wait for the write to finish. Another write may start after this one finishes, unless the caller holds the folio lock.

int **folio_wait_fscache_killable**(struct *folio* *folio)

Wait for an fscache write on this folio to end.

**Parameters**

**struct folio *folio**

The folio.

**Description**

If this folio is currently being written to a local cache, wait for the write to finish or for a fatal signal to be received. Another write may start after this one finishes, unless the caller holds the folio lock.

**Return**

- 0 if successful.
- -EINTR if a fatal signal was encountered.

struct *netfs_inode* ***netfs_inode**(struct *inode* *inode)

Get the netfs inode context from the inode

**Parameters**

**struct inode *inode**

The inode to query

**Description**

Get the netfs lib inode context from the network filesystem's inode. The context struct is expected to directly follow on from the VFS inode struct.

void **netfs_inode_init**(struct *netfs_inode* *ctx, const struct netfs_request_ops *ops, bool use_zero_point)

Initialise a netfslib inode context

**Parameters**

**struct netfs_inode *ctx**

The netfs inode to initialise

**const struct netfs_request_ops *ops**

The netfs's operations list

**bool use_zero_point**

True to use the zero_point read optimisation

**Description**

Initialise the netfs library context struct. This is expected to follow on directly from the VFS inode struct.

void **netfs_resize_file**(struct *netfs_inode* *ctx, loff_t new_i_size, bool changed_on_server)

Note that a file got resized

**Parameters**

**struct netfs_inode *ctx**

The netfs inode being resized

**loff_t new_i_size**

The new file size

**bool changed_on_server**

The change was applied to the server

**Description**

Inform the netfs lib that a file got resized so that it can adjust its state.

struct fscache_cookie ***netfs_i_cookie**(struct *netfs_inode* *ctx)

Get the cache cookie from the inode

**Parameters**

**struct netfs_inode *ctx**

The netfs inode to query

**Description**

Get the caching cookie (if enabled) from the network filesystem's inode.

void **netfs_readahead**(struct readahead_control *ractl)

Helper to manage a read request

**Parameters**

**struct readahead_control *ractl**

The description of the readahead request

**Description**

Fulfil a readahead request by drawing data from the cache if possible, or the netfs if not. Space beyond the EOF is zero-filled. Multiple I/O requests from different sources will get munged together. If necessary, the readahead window can be expanded in either direction to a more convenient alighment for RPC efficiency or to make storage in the cache feasible.

The calling netfs must initialise a netfs context contiguous to the vfs inode before calling this.

This is usable whether or not caching is enabled.

int **netfs_read_folio**(struct *file* *file, struct *folio* *folio)

Helper to manage a read_folio request

**Parameters**

**struct file *file**

The file to read from

**struct folio *folio**
The folio to read

**Description**

Fulfil a read_folio request by drawing data from the cache if possible, or the netfs if not. Space beyond the EOF is zero-filled. Multiple I/O requests from different sources will get munged together.

The calling netfs must initialise a netfs context contiguous to the vfs inode before calling this.

This is usable whether or not caching is enabled.

int **netfs_write_begin**(struct *netfs_inode* *ctx, struct *file* *file, struct *address_space* *mapping, loff_t pos, unsigned int len, struct folio **_folio, void **_fsdata)

Helper to prepare for writing

**Parameters**

**struct netfs_inode *ctx**
The netfs context

**struct file *file**
The file to read from

**struct address_space *mapping**
The mapping to read from

**loff_t pos**
File position at which the write will begin

**unsigned int len**
The length of the write (may extend beyond the end of the folio chosen)

**struct folio **_folio**
Where to put the resultant folio

**void **_fsdata**
Place for the netfs to store a cookie

**Description**

Pre-read data for a write-begin request by drawing data from the cache if possible, or the netfs if not. Space beyond the EOF is zero-filled. Multiple I/O requests from different sources will get munged together. If necessary, the readahead window can be expanded in either direction to a more convenient alighment for RPC efficiency or to make storage in the cache feasible.

The calling netfs must provide a table of operations, only one of which, issue_op, is mandatory.

The check_write_begin() operation can be provided to check for and flush conflicting writes once the folio is grabbed and locked. It is passed a pointer to the fsdata cookie that gets returned to the VM to be passed to write_end. It is permitted to sleep. It should return 0 if the request should go ahead or it may return an error. It may also unlock and put the folio, provided it sets `*foliop` to NULL, in which case a return of 0 will cause the folio to be re-got and the process to be retried.

The calling netfs must initialise a netfs context contiguous to the vfs inode before calling this.

This is usable whether or not caching is enabled.

ssize_t **netfs_buffered_read_iter**(struct kiocb *iocb, struct iov_iter *iter)

Filesystem buffered I/O read routine

**Parameters**

**struct kiocb *iocb**
kernel I/O control block

**struct iov_iter *iter**
destination for the data read

**Description**

This is the ->read_iter() routine for all filesystems that can use the page cache directly.

The IOCB_NOWAIT flag in iocb->ki_flags indicates that -EAGAIN shall be returned when no data can be read without waiting for I/O requests to complete; it doesn't prevent readahead.

The IOCB_NOIO flag in iocb->ki_flags indicates that no new I/O requests shall be made for the read or for readahead. When no data can be read, -EAGAIN shall be returned. When readahead would be triggered, a partial, possibly empty read shall be returned.

**Return**

- number of bytes copied, even for partial reads
- negative error code (or 0 if IOCB_NOIO) if nothing was read

ssize_t **netfs_file_read_iter**(struct kiocb *iocb, struct iov_iter *iter)

Generic filesystem read routine

**Parameters**

**struct kiocb *iocb**
kernel I/O control block

**struct iov_iter *iter**
destination for the data read

**Description**

This is the ->read_iter() routine for all filesystems that can use the page cache directly.

The IOCB_NOWAIT flag in iocb->ki_flags indicates that -EAGAIN shall be returned when no data can be read without waiting for I/O requests to complete; it doesn't prevent readahead.

The IOCB_NOIO flag in iocb->ki_flags indicates that no new I/O requests shall be made for the read or for readahead. When no data can be read, -EAGAIN shall be returned. When readahead would be triggered, a partial, possibly empty read shall be returned.

**Return**

- number of bytes copied, even for partial reads
- negative error code (or 0 if IOCB_NOIO) if nothing was read

void **netfs_subreq_terminated**(struct netfs_io_subrequest *subreq, ssize_t transferred_or_error, bool was_async)

Note the termination of an I/O operation.

**Parameters**

**struct netfs_io_subrequest *subreq**
The I/O request that has terminated.

**ssize_t transferred_or_error**
The amount of data transferred or an error code.

**bool was_async**
The termination was asynchronous

**Description**

This tells the read helper that a contributory I/O operation has terminated, one way or another, and that it should integrate the results.

The caller indicates in **transferred_or_error** the outcome of the operation, supplying a positive value to indicate the number of bytes transferred, 0 to indicate a failure to transfer anything that should be retried or a negative error code. The helper will look after reissuing I/O operations as appropriate and writing downloaded data to the cache.

If **was_async** is true, the caller might be running in softirq or interrupt context and we can't sleep.

CHAPTER

THREE

# FILESYSTEMS

Documentation for filesystem implementations.

## 3.1 v9fs: Plan 9 Resource Sharing for Linux

### 3.1.1 About

v9fs is a Unix implementation of the Plan 9 9p remote filesystem protocol.

This software was originally developed by Ron Minnich <rminnich@sandia.gov> and Maya Gokhale. Additional development by Greg Watson <gwatson@lanl.gov> and most recently Eric Van Hensbergen <ericvh@gmail.com>, Latchesar Ionkov <lucho@ionkov.net> and Russ Cox <rsc@swtch.com>.

The best detailed explanation of the Linux implementation and applications of the 9p client is available in the form of a USENIX paper:

> https://www.usenix.org/events/usenix05/tech/freenix/hensbergen.html

Other applications are described in the following papers:

- XCPU & Clustering http://xcpu.org/papers/xcpu-talk.pdf
- KVMFS: control file system for KVM http://xcpu.org/papers/kvmfs.pdf
- CellFS: A New Programming Model for the Cell BE http://xcpu.org/papers/cellfs-talk.pdf
- PROSE I/O: Using 9p to enable Application Partitions http://plan9.escet.urjc.es/iwp9/cready/PROSE_iwp9_2006.pdf
- VirtFS: A Virtualization Aware File System pass-through http://goo.gl/3WPDg

### 3.1.2 Usage

For remote file server:

```
mount -t 9p 10.10.1.2 /mnt/9
```

For Plan 9 From User Space applications (http://swtch.com/plan9):

```
mount -t 9p `namespace`/acme /mnt/9 -o trans=unix,uname=$USER
```

For server running on QEMU host with virtio transport:

```
mount -t 9p -o trans=virtio <mount_tag> /mnt/9
```

where mount_tag is the tag associated by the server to each of the exported mount points. Each 9P export is seen by the client as a virtio device with an associated ”mount_tag” property. Available mount tags can be seen by reading /sys/bus/virtio/drivers/9pnet_virtio/virtio<n>/mount_tag files.

### 3.1.3 Options

<table>
<tr><td>trans=name</td><td>select an alternative transport. Valid options are currently:<table><tr><td>unix</td><td>specifying a named pipe mount point</td></tr><tr><td>tcp</td><td>specifying a normal TCP/IP connection</td></tr><tr><td>fd</td><td>used passed file descriptors for connection (see rfdno and wfdno)</td></tr><tr><td>virtio</td><td>connect to the next virtio channel available (from QEMU with trans_virtio module)</td></tr><tr><td>rdma</td><td>connect to a specified RDMA channel</td></tr></table></td></tr>
<tr><td>uname=name</td><td>user name to attempt mount as on the remote server. The server may override or ignore this value. Certain user names may require authentication.</td></tr>
<tr><td>aname=name</td><td>aname specifies the file tree to access when the server is offering several exported file systems.</td></tr>
<tr><td>cache=mode</td><td>specifies a caching policy. By default, no caches are used. The mode can be specified as a bitmask or by using one of the preexisting common 'shortcuts'. The bitmask is described below: (unspecified bits are reserved)<table><tr><td>0b00000000</td><td>all caches disabled, mmap disabled</td></tr><tr><td>0b00000001</td><td>file caches enabled</td></tr><tr><td>0b00000010</td><td>meta-data caches enabled</td></tr><tr><td>0b00000100</td><td>writeback behavior (as opposed to writethrough)</td></tr><tr><td>0b00001000</td><td>loose caches (no explicit consistency with server)</td></tr><tr><td>0b10000000</td><td>fscache enabled for persistent caching</td></tr></table></td></tr>
</table>

### 3.1.4 Behavior

This section aims at describing 9p 'quirks' that can be different from a local filesystem behaviors.

- Setting O_NONBLOCK on a file will make client reads return as early as the server returns some data instead of trying to fill the read buffer with the requested amount of bytes or end of file is reached.

### 3.1.5 Resources

Protocol specifications are maintained on github: http://ericvh.github.com/9p-rfc/

9p client and server implementations are listed on http://9p.cat-v.org/implementations

A 9p2000.L server is being developed by LLNL and can be found at http://code.google.com/p/diod/

There are user and developer mailing lists available through the v9fs project on sourceforge (http://sourceforge.net/projects/v9fs).

News and other information is maintained on a Wiki. (http://sf.net/apps/mediawiki/v9fs/index.php).

Bug reports are best issued via the mailing list.

For more information on the Plan 9 Operating System check out http://plan9.bell-labs.com/plan9

For information on Plan 9 from User Space (Plan 9 applications and libraries ported to Linux/BSD/OSX/etc) check out https://9fans.github.io/plan9port/

## 3.2 Acorn Disc Filing System - ADFS

### 3.2.1 Filesystems supported by ADFS

The ADFS module supports the following Filecore formats which have:

- new maps
- new directories or big directories

In terms of the named formats, this means we support:

- E and E+, with or without boot block
- F and F+

We fully support reading files from these filesystems, and writing to existing files within their existing allocation. Essentially, we do not support changing any of the filesystem metadata.

This is intended to support loopback mounted Linux native filesystems on a RISC OS Filecore filesystem, but will allow the data within files to be changed.

If write support (ADFS_FS_RW) is configured, we allow rudimentary directory updates, specifically updating the access mode and timestamp.

### 3.2.2 Mount options for ADFS

| | |
|---|---|
| uid=nnn | All files in the partition will be owned by user id nnn. Default 0 (root). |
| gid=nnn | All files in the partition will be in group nnn. Default 0 (root). |
| ownmask=nnn | The permission mask for ADFS 'owner' permissions will be nnn. Default 0700. |
| othmask=nnn | The permission mask for ADFS 'other' permissions will be nnn. Default 0077. |
| ftsuffix=n | When ftsuffix=0, no file type suffix will be applied. When ftsuffix=1, a hexadecimal suffix corresponding to the RISC OS file type will be added. Default 0. |

### 3.2.3 Mapping of ADFS permissions to Linux permissions

ADFS permissions consist of the following:

- Owner read
- Owner write
- Other read
- Other write

(In older versions, an 'execute' permission did exist, but this does not hold the same meaning as the Linux 'execute' permission and is now obsolete).

The mapping is performed as follows:

```
    Owner read                              -> -r--r--r--
    Owner write                             -> --w--w---w
    Owner read and filetype UnixExec        -> ---x--x--x
These are then masked by ownmask, eg 700    -> -rwx------
    Possible owner mode permissions         -> -rwx------

    Other read                              -> -r--r--r--
    Other write                             -> --w--w--w-
    Other read and filetype UnixExec        -> ---x--x--x
These are then masked by othmask, eg 077    -> ----rwxrwx
    Possible other mode permissions         -> ----rwxrwx
```

Hence, with the default masks, if a file is owner read/write, and not a UnixExec filetype, then the permissions will be:

```
-rw-------
```

However, if the masks were ownmask=0770,othmask=0007, then this would be modified to:

```
-rw-rw----
```

There is no restriction on what you can do with these masks. You may wish that either read bits give read access to the file for all, but keep the default write protection (ownmask=0755,othmask=0577):

```
-rw-r--r--
```

You can therefore tailor the permission translation to whatever you desire the permissions should be under Linux.

### 3.2.4 RISC OS file type suffix

RISC OS file types are stored in bits 19..8 of the file load address.

To enable non-RISC OS systems to be used to store files without losing file type information, a file naming convention was devised (initially for use with NFS) such that a hexadecimal suffix of the form ,xyz denoted the file type: e.g. BasicFile,ffb is a BASIC (0xffb) file. This naming convention is now also used by RISC OS emulators such as RPCEmu.

Mounting an ADFS disc with option ftsuffix=1 will cause appropriate file type suffixes to be appended to file names read from a directory. If the ftsuffix option is zero or omitted, no file type suffixes will be added.

## 3.3 Overview of Amiga Filesystems

Not all varieties of the Amiga filesystems are supported for reading and writing. The Amiga currently knows six different filesystems:

| | |
|---|---|
| DOS0 | The old or original filesystem, not really suited for hard disks and normally not used on them, either. Supported read/write. |
| DOS1 | The original Fast File System. Supported read/write. |
| DOS2 | The old ”international” filesystem. International means that a bug has been fixed so that accented (”international”) letters in file names are case-insensitive, as they ought to be. Supported read/write. |
| DOS3 | The ”international” Fast File System. Supported read/write. |
| DOS4 | The original filesystem with directory cache. The directory cache speeds up directory accesses on floppies considerably, but slows down file creation/deletion. Doesn't make much sense on hard disks. Supported read only. |
| DOS5 | The Fast File System with directory cache. Supported read only. |

All of the above filesystems allow block sizes from 512 to 32K bytes. Supported block sizes are: 512, 1024, 2048 and 4096 bytes. Larger blocks speed up almost everything at the expense of wasted disk space. The speed gain above 4K seems not really worth the price, so you don't lose too much here, either.

The muFS (multi user File System) equivalents of the above file systems are supported, too.

### 3.3.1 Mount options for the AFFS

**protect**
If this option is set, the protection bits cannot be altered.

**setuid[=uid]**
This sets the owner of all files and directories in the file system to uid or the uid of the current user, respectively.

**setgid[=gid]**
Same as above, but for gid.

**mode=mode**
Sets the mode flags to the given (octal) value, regardless of the original permissions. Directories will get an x permission if the corresponding r bit is set. This is useful since most of the plain AmigaOS files will map to 600.

**nofilenametruncate**
The file system will return an error when filename exceeds standard maximum filename length (30 characters).

**reserved=num**
Sets the number of reserved blocks at the start of the partition to num. You should never need this option. Default is 2.

**root=block**
Sets the block number of the root block. This should never be necessary.

**bs=blksize**
Sets the blocksize to blksize. Valid block sizes are 512, 1024, 2048 and 4096. Like the root option, this should never be necessary, as the affs can figure it out itself.

**quiet**
The file system will not return an error for disallowed mode changes.

**verbose**
The volume name, file system type and block size will be written to the syslog when the filesystem is mounted.

**mufs**
The filesystem is really a muFS, also it doesn't identify itself as one. This option is necessary if the filesystem wasn't formatted as muFS, but is used as one.

**prefix=path**
Path will be prefixed to every absolute path name of symbolic links on an AFFS partition. Default = ”/”. (See below.)

**volume=name**
When symbolic links with an absolute path are created on an AFFS partition, name will be prepended as the volume name. Default = ”” (empty string). (See below.)

### 3.3.2 Handling of the Users/Groups and protection flags

Amiga -> Linux:

The Amiga protection flags RWEDRWEDHSPARWED are handled as follows:

- R maps to r for user, group and others. On directories, R implies x.
- W maps to w.
- E maps to x.
- D is ignored.
- H, S and P are always retained and ignored under Linux.
- A is cleared when a file is written to.

User id and group id will be used unless set[gu]id are given as mount options. Since most of the Amiga file systems are single user systems they will be owned by root. The root directory (the mount point) of the Amiga filesystem will be owned by the user who actually mounts the filesystem (the root directory doesn't have uid/gid fields).

Linux -> Amiga:

The Linux rwxrwxrwx file mode is handled as follows:

- r permission will allow R for user, group and others.
- w permission will allow W for user, group and others.
- x permission of the user will allow E for plain files.
- D will be allowed for user, group and others.
- All other flags (suid, sgid, ...) are ignored and will not be retained.

Newly created files and directories will get the user and group ID of the current user and a mode according to the umask.

### 3.3.3 Symbolic links

Although the Amiga and Linux file systems resemble each other, there are some, not always subtle, differences. One of them becomes apparent with symbolic links. While Linux has a file system with exactly one root directory, the Amiga has a separate root directory for each file system (for example, partition, floppy disk, ...). With the Amiga, these entities are called ”volumes”. They have symbolic names which can be used to access them. Thus, symbolic links can point to a different volume. AFFS turns the volume name into a directory name and prepends the prefix path (see prefix option) to it.

Example: You mount all your Amiga partitions under /amiga/<volume> (where <volume> is the name of the volume), and you give the option ”prefix=/amiga/” when mounting all your AFFS partitions. (They might be ”User”, ”WB” and ”Graphics”, the mount points /amiga/User, /amiga/WB and /amiga/Graphics). A symbolic link referring to ”User:sc/include/dos/dos.h” will be followed to ”/amiga/User/sc/include/dos/dos.h”.

### 3.3.4 Examples

Command line:

```
mount  Archive/Amiga/Workbench3.1.adf /mnt -t affs -o loop,verbose
mount  /dev/sda3 /Amiga -t affs
```

/etc/fstab entry:

```
/dev/sdb5   /amiga/Workbench    affs    noauto,user,exec,verbose 0 0
```

### 3.3.5 IMPORTANT NOTE

If you boot Windows 95 (don't know about 3.x, 98 and NT) while you have an Amiga harddisk connected to your PC, it will overwrite the bytes 0x00dc..0x00df of block 0 with garbage, thus invalidating the Rigid Disk Block. Sheer luck has it that this is an unused area of the RDB, so only the checksum doesn't match anymore. Linux will ignore this garbage and recognize the RDB anyway, but before you connect that drive to your Amiga again, you must restore or repair your RDB. So please do make a backup copy of it before booting Windows!

If the damage is already done, the following should fix the RDB (where <disk> is the device name).

DO AT YOUR OWN RISK:

```
dd if=/dev/<disk> of=rdb.tmp count=1
cp rdb.tmp rdb.fixed
dd if=/dev/zero of=rdb.fixed bs=1 seek=220 count=4
dd if=rdb.fixed of=/dev/<disk>
```

### 3.3.6 Bugs, Restrictions, Caveats

Quite a few things may not work as advertised. Not everything is tested, though several hundred MB have been read and written using this fs. For a most up-to-date list of bugs please consult fs/affs/Changes.

By default, filenames are truncated to 30 characters without warning. 'nofilenametruncate' mount option can change that behavior.

Case is ignored by the affs in filename matching, but Linux shells do care about the case. Example (with /wb being an affs mounted fs):

```
rm /wb/WRONGCASE
```

will remove /mnt/wrongcase, but:

```
rm /wb/WR*
```

will not since the names are matched by the shell.

The block allocation is designed for hard disk partitions. If more than 1 process writes to a (small) diskette, the blocks are allocated in an ugly way (but the real AFFS doesn't do much better). This is also true when space gets tight.

You cannot execute programs on an OFS (Old File System), since the program files cannot be memory mapped due to the 488 byte blocks. For the same reason you cannot mount an image on such a filesystem via the loopback device.

The bitmap valid flag in the root block may not be accurate when the system crashes while an affs partition is mounted. There's currently no way to fix a garbled filesystem without an Amiga (disk validator) or manually (who would do this?). Maybe later.

If you mount affs partitions on system startup, you may want to tell fsck that the fs should not be checked (place a '0' in the sixth field of /etc/fstab).

It's not possible to read floppy disks with a normal PC or workstation due to an incompatibility with the Amiga floppy controller.

If you are interested in an Amiga Emulator for Linux, look at

http://web.archive.org/web/%2E/http://www.freiburg.linux.de/~uae/

## 3.4 kAFS: AFS FILESYSTEM

### 3.4.1 Overview

This filesystem provides a fairly simple secure AFS filesystem driver. It is under development and does not yet provide the full feature set. The features it does support include:

(*) Security (currently only AFS kaserver and KerberosIV tickets).

(*) File reading and writing.

(*) Automounting.

(*) Local caching (via fscache).

It does not yet support the following AFS features:

(*) pioctl() system call.

### 3.4.2 Compilation

The filesystem should be enabled by turning on the kernel configuration options:

```
CONFIG_AF_RXRPC         - The RxRPC protocol transport
CONFIG_RXKAD            - The RxRPC Kerberos security handler
CONFIG_AFS_FS           - The AFS filesystem
```

Additionally, the following can be turned on to aid debugging:

```
CONFIG_AF_RXRPC_DEBUG   - Permit AF_RXRPC debugging to be enabled
CONFIG_AFS_DEBUG        - Permit AFS debugging to be enabled
```

They permit the debugging messages to be turned on dynamically by manipulating the masks in the following files:

```
/sys/module/af_rxrpc/parameters/debug
/sys/module/kafs/parameters/debug
```

### 3.4.3 Usage

When inserting the driver modules the root cell must be specified along with a list of volume location server IP addresses:

```
modprobe rxrpc
modprobe kafs rootcell=cambridge.redhat.com:172.16.18.73:172.16.18.91
```

The first module is the AF_RXRPC network protocol driver. This provides the RxRPC remote operation protocol and may also be accessed from userspace. See:

Documentation/networking/rxrpc.rst

The second module is the kerberos RxRPC security driver, and the third module is the actual filesystem driver for the AFS filesystem.

Once the module has been loaded, more modules can be added by the following procedure:

```
echo add grand.central.org 18.9.48.14:128.2.203.61:130.237.48.87 >/proc/fs/afs/
↪cells
```

Where the parameters to the "add" command are the name of a cell and a list of volume location servers within that cell, with the latter separated by colons.

Filesystems can be mounted anywhere by commands similar to the following:

```
mount -t afs "%cambridge.redhat.com:root.afs." /afs
mount -t afs "#cambridge.redhat.com:root.cell." /afs/cambridge
mount -t afs "#root.afs." /afs
mount -t afs "#root.cell." /afs/cambridge
```

Where the initial character is either a hash or a percent symbol depending on whether you definitely want a R/W volume (percent) or whether you'd prefer a R/O volume, but are willing to use a R/W volume instead (hash).

The name of the volume can be suffixes with ".backup" or ".readonly" to specify connection to only volumes of those types.

The name of the cell is optional, and if not given during a mount, then the named volume will be looked up in the cell specified during modprobe.

Additional cells can be added through /proc (see later section).

### 3.4.4 Mountpoints

AFS has a concept of mountpoints. In AFS terms, these are specially formatted symbolic links (of the same form as the "device name" passed to mount). kAFS presents these to the user as directories that have a follow-link capability (i.e.: symbolic link semantics). If anyone attempts to access them, they will automatically cause the target volume to be mounted (if possible) on that site.

Automatically mounted filesystems will be automatically unmounted approximately twenty minutes after they were last used. Alternatively they can be unmounted directly with the umount() system call.

Manually unmounting an AFS volume will cause any idle submounts upon it to be culled first. If all are culled, then the requested volume will also be unmounted, otherwise error EBUSY will be returned.

This can be used by the administrator to attempt to unmount the whole AFS tree mounted on /afs in one go by doing:

```
umount /afs
```

### 3.4.5 Dynamic Root

A mount option is available to create a serverless mount that is only usable for dynamic lookup. Creating such a mount can be done by, for example:

```
mount -t afs none /afs -o dyn
```

This creates a mount that just has an empty directory at the root. Attempting to look up a name in this directory will cause a mountpoint to be created that looks up a cell of the same name, for example:

```
ls /afs/grand.central.org/
```

### 3.4.6 Proc Filesystem

The AFS module creates a ”/proc/fs/afs/” directory and populates it:

**(*) A ”cells” file that lists cells currently known to the afs module and** their usage counts:

```
[root@andromeda ~]# cat /proc/fs/afs/cells
USE NAME
  3 cambridge.redhat.com
```

**(*) A directory per cell that contains files that list volume location** servers, volumes, and active servers known within that cell:

```
[root@andromeda ~]# cat /proc/fs/afs/cambridge.redhat.com/servers
USE ADDR            STATE
  4 172.16.18.91        0
[root@andromeda ~]# cat /proc/fs/afs/cambridge.redhat.com/vlservers
ADDRESS
172.16.18.91
[root@andromeda ~]# cat /proc/fs/afs/cambridge.redhat.com/volumes
USE STT VLID[0]  VLID[1]  VLID[2]  NAME
  1 Val 20000000 20000001 20000002 root.afs
```

### 3.4.7 The Cell Database

The filesystem maintains an internal database of all the cells it knows and the IP addresses of the volume location servers for those cells. The cell to which the system belongs is added to the database when modprobe is performed by the "rootcell=" argument or, if compiled in, using a "kafs.rootcell=" argument on the kernel command line.

Further cells can be added by commands similar to the following:

```
echo add CELLNAME VLADDR[:VLADDR][:VLADDR]... >/proc/fs/afs/cells
echo add grand.central.org 18.9.48.14:128.2.203.61:130.237.48.87 >/proc/fs/afs/
↪cells
```

No other cell database operations are available at this time.

### 3.4.8 Security

Secure operations are initiated by acquiring a key using the klog program. A very primitive klog program is available at:

https://people.redhat.com/~dhowells/rxrpc/klog.c

This should be compiled by:

```
make klog LDLIBS="-lcrypto -lcrypt -lkrb4 -lkeyutils"
```

And then run as:

```
./klog
```

Assuming it's successful, this adds a key of type RxRPC, named for the service and cell, e.g.: "afs@<cellname>". This can be viewed with the keyctl program or by cat'ing /proc/keys:

```
[root@andromeda ~]# keyctl show
Session Keyring
       -3 --alswrv      0     0  keyring: _ses.3268
        2 --alswrv      0     0   \_ keyring: _uid.0
111416553 --als--v      0     0   \_ rxrpc: afs@CAMBRIDGE.REDHAT.COM
```

Currently the username, realm, password and proposed ticket lifetime are compiled into the program.

It is not required to acquire a key before using AFS facilities, but if one is not acquired then all operations will be governed by the anonymous user parts of the ACLs.

If a key is acquired, then all AFS operations, including mounts and automounts, made by a possessor of that key will be secured with that key.

If a file is opened with a particular key and then the file descriptor is passed to a process that doesn't have that key (perhaps over an AF_UNIX socket), then the operations on the file will be made with key that was used to open the file.

### 3.4.9 The @sys Substitution

The list of up to 16 @sys substitutions for the current network namespace can be configured by writing a list to /proc/fs/afs/sysname:

```
[root@andromeda ~]# echo foo amd64_linux_26 >/proc/fs/afs/sysname
```

or cleared entirely by writing an empty list:

```
[root@andromeda ~]# echo >/proc/fs/afs/sysname
```

The current list for current network namespace can be retrieved by:

```
[root@andromeda ~]# cat /proc/fs/afs/sysname
foo
amd64_linux_26
```

When @sys is being substituted for, each element of the list is tried in the order given.

By default, the list will contain one item that conforms to the pattern "<arch>_linux_26", amd64 being the name for x86_64.

## 3.5 autofs - how it works

### 3.5.1 Purpose

The goal of autofs is to provide on-demand mounting and race free automatic unmounting of various other filesystems. This provides two key advantages:

1. There is no need to delay boot until all filesystems that might be needed are mounted. Processes that try to access those slow filesystems might be delayed but other processes can continue freely. This is particularly important for network filesystems (e.g. NFS) or filesystems stored on media with a media-changing robot.

2. The names and locations of filesystems can be stored in a remote database and can change at any time. The content in that data base at the time of access will be used to provide a target for the access. The interpretation of names in the filesystem can even be programmatic rather than database-backed, allowing wildcards for example, and can vary based on the user who first accessed a name.

### 3.5.2 Context

The "autofs" filesystem module is only one part of an autofs system. There also needs to be a user-space program which looks up names and mounts filesystems. This will often be the "automount" program, though other tools including "systemd" can make use of "autofs". This document describes only the kernel module and the interactions required with any user-space program. Subsequent text refers to this as the "automount daemon" or simply "the daemon".

"autofs" is a Linux kernel module which provides the "autofs" filesystem type. Several "autofs" filesystems can be mounted and they can each be managed separately, or all managed by the same daemon.

### 3.5.3 Content

An autofs filesystem can contain 3 sorts of objects: directories, symbolic links and mount traps. Mount traps are directories with extra properties as described in the next section.

Objects can only be created by the automount daemon: symlinks are created with a regular *symlink* system call, while directories and mount traps are created with *mkdir*. The determination of whether a directory should be a mount trap is based on a master map. This master map is consulted by autofs to determine which directories are mount points. Mount points can be *direct*/*indirect*/*offset*. On most systems, the default master map is located at */etc/auto.master*.

If neither the *direct* or *offset* mount options are given (so the mount is considered to be *indirect*), then the root directory is always a regular directory, otherwise it is a mount trap when it is empty and a regular directory when not empty. Note that *direct* and *offset* are treated identically so a concise summary is that the root directory is a mount trap only if the filesystem is mounted *direct* and the root is empty.

Directories created in the root directory are mount traps only if the filesystem is mounted *indirect* and they are empty.

Directories further down the tree depend on the *maxproto* mount option and particularly whether it is less than five or not. When *maxproto* is five, no directories further down the tree are ever mount traps, they are always regular directories. When the *maxproto* is four (or three), these directories are mount traps precisely when they are empty.

So: non-empty (i.e. non-leaf) directories are never mount traps. Empty directories are sometimes mount traps, and sometimes not depending on where in the tree they are (root, top level, or lower), the *maxproto*, and whether the mount was *indirect* or not.

### 3.5.4 Mount Traps

A core element of the implementation of autofs is the Mount Traps which are provided by the Linux VFS. Any directory provided by a filesystem can be designated as a trap. This involves two separate features that work together to allow autofs to do its job.

**DCACHE_NEED_AUTOMOUNT**

If a dentry has the DCACHE_NEED_AUTOMOUNT flag set (which gets set if the inode has S_AUTOMOUNT set, or can be set directly) then it is (potentially) a mount trap. Any access to this directory beyond a "*stat*" will (normally) cause the *d_op->d_automount()* dentry operation to be called. The task of this method is to find the filesystem that should be mounted on the directory and to return it. The VFS is responsible for actually mounting the root of this filesystem on the directory.

autofs doesn't find the filesystem itself but sends a message to the automount daemon asking it to find and mount the filesystem. The autofs *d_automount* method then waits for the daemon to report that everything is ready. It will then return "*NULL*" indicating that the mount has already happened. The VFS doesn't try to mount anything but follows down the mount that is already there.

This functionality is sufficient for some users of mount traps such as NFS which creates traps so that mountpoints on the server can be reflected on the client. However it is not sufficient for autofs. As mounting onto a directory is considered to be "beyond a *stat*", the automount daemon would not be able to mount a filesystem on the 'trap' directory without some way to avoid getting caught in the trap. For that purpose there is another flag.

**DCACHE_MANAGE_TRANSIT**

If a dentry has DCACHE_MANAGE_TRANSIT set then two very different but related behaviours are invoked, both using the *d_op->d_manage()* dentry operation.

Firstly, before checking to see if any filesystem is mounted on the directory, d_manage() will be called with the *rcu_walk* parameter set to *false*. It may return one of three things:

- A return value of zero indicates that there is nothing special about this dentry and normal checks for mounts and automounts should proceed.

  autofs normally returns zero, but first waits for any expiry (automatic unmounting of the mounted filesystem) to complete. This avoids races.

- A return value of *-EISDIR* tells the VFS to ignore any mounts on the directory and to not consider calling *->d_automount()*. This effectively disables the **DCACHE_NEED_AUTOMOUNT** flag causing the directory not be a mount trap after all.

  autofs returns this if it detects that the process performing the lookup is the automount daemon and that the mount has been requested but has not yet completed. How it determines this is discussed later. This allows the automount daemon not to get caught in the mount trap.

  There is a subtlety here. It is possible that a second autofs filesystem can be mounted below the first and for both of them to be managed by the same daemon. For the daemon to be able to mount something on the second it must be able to ”walk” down past the first. This means that d_manage cannot *always* return -EISDIR for the automount daemon. It must only return it when a mount has been requested, but has not yet completed.

  *d_manage* also returns *-EISDIR* if the dentry shouldn't be a mount trap, either because it is a symbolic link or because it is not empty.

- Any other negative value is treated as an error and returned to the caller.

  autofs can return

  - -ENOENT if the automount daemon failed to mount anything,
  - -ENOMEM if it ran out of memory,
  - -EINTR if a signal arrived while waiting for expiry to complete
  - or any other error sent down by the automount daemon.

The second use case only occurs during an ”RCU-walk” and so *rcu_walk* will be set.

An RCU-walk is a fast and lightweight process for walking down a filename path (i.e. it is like running on tip-toes). RCU-walk cannot cope with all situations so when it finds a difficulty it falls back to ”REF-walk”, which is slower but more robust.

RCU-walk will never call *->d_automount*; the filesystems must already be mounted or RCU-walk cannot handle the path. To determine if a mount-trap is safe for RCU-walk mode it calls *->d_manage()* with *rcu_walk* set to *true*.

In this case *d_manage()* must avoid blocking and should avoid taking spinlocks if at all possible. Its sole purpose is to determine if it would be safe to follow down into any mounted directory and the only reason that it might not be is if an expiry of the mount is underway.

In the *rcu_walk* case, *d_manage()* cannot return -EISDIR to tell the VFS that this is a directory that doesn't require d_automount. If *rcu_walk* sees a dentry with DCACHE_NEED_AUTOMOUNT set but nothing mounted, it *will* fall back to REF-walk.

*d_manage()* cannot make the VFS remain in RCU-walk mode, but can only tell it to get out of RCU-walk mode by returning *-ECHILD*.

So *d_manage()*, when called with *rcu_walk* set, should either return -ECHILD if there is any reason to believe it is unsafe to enter the mounted filesystem, otherwise it should return 0.

autofs will return *-ECHILD* if an expiry of the filesystem has been initiated or is being considered, otherwise it returns 0.

### 3.5.5 Mountpoint expiry

The VFS has a mechanism for automatically expiring unused mounts, much as it can expire any unused dentry information from the dcache. This is guided by the MNT_SHRINKABLE flag. This only applies to mounts that were created by *d_automount()* returning a filesystem to be mounted. As autofs doesn't return such a filesystem but leaves the mounting to the automount daemon, it must involve the automount daemon in unmounting as well. This also means that autofs has more control over expiry.

The VFS also supports ”expiry” of mounts using the MNT_EXPIRE flag to the *umount* system call. Unmounting with MNT_EXPIRE will fail unless a previous attempt had been made, and the filesystem has been inactive and untouched since that previous attempt. autofs does not depend on this but has its own internal tracking of whether filesystems were recently used. This allows individual names in the autofs directory to expire separately.

With version 4 of the protocol, the automount daemon can try to unmount any filesystems mounted on the autofs filesystem or remove any symbolic links or empty directories any time it likes. If the unmount or removal is successful the filesystem will be returned to the state it was before the mount or creation, so that any access of the name will trigger normal auto-mount processing. In particular, *rmdir* and *unlink* do not leave negative entries in the dcache as a normal filesystem would, so an attempt to access a recently-removed object is passed to autofs for handling.

With version 5, this is not safe except for unmounting from top-level directories. As lower-level directories are never mount traps, other processes will see an empty directory as soon as the filesystem is unmounted. So it is generally safest to use the autofs expiry protocol described below.

Normally the daemon only wants to remove entries which haven't been used for a while. For this purpose autofs maintains a ”*last_used*” time stamp on each directory or symlink. For symlinks it genuinely does record the last time the symlink was ”used” or followed to find out where it points to. For directories the field is used slightly differently. The field is updated at mount time and during expire checks if it is found to be in use (ie. open file descriptor or process working directory) and during path walks. The update done during path walks prevents frequent expire and immediate mount of frequently accessed automounts. But in the case where a GUI continually access or an application frequently scans an autofs directory tree there can be an accumulation of mounts that aren't actually being used. To cater for this case the ”*strictexpire*” autofs mount option can be used to avoid the ”*last_used*” update on path walk thereby preventing this apparent inability to expire mounts that aren't really in use.

The daemon is able to ask autofs if anything is due to be expired, using an *ioctl* as discussed later. For a *direct* mount, autofs considers if the entire mount-tree can be unmounted or not. For an *indirect* mount, autofs considers each of the names in the top level directory to determine if any of those can be unmounted and cleaned up.

There is an option with indirect mounts to consider each of the leaves that has been mounted on instead of considering the top-level names. This was originally intended for compatibility with version 4 of autofs and should be considered as deprecated for Sun Format automount maps. However, it may be used again for amd format mount maps (which are generally indirect maps) because the amd automounter allows for the setting of an expire timeout for individual mounts. But there are some difficulties in making the needed changes for this.

When autofs considers a directory it checks the *last_used* time and compares it with the "timeout" value set when the filesystem was mounted, though this check is ignored in some cases. It also checks if the directory or anything below it is in use. For symbolic links, only the *last_used* time is ever considered.

If both appear to support expiring the directory or symlink, an action is taken.

There are two ways to ask autofs to consider expiry. The first is to use the **AUTOFS_IOC_EXPIRE** ioctl. This only works for indirect mounts. If it finds something in the root directory to expire it will return the name of that thing. Once a name has been returned the automount daemon needs to unmount any filesystems mounted below the name normally. As described above, this is unsafe for non-toplevel mounts in a version-5 autofs. For this reason the current *automount(8)* does not use this ioctl.

The second mechanism uses either the **AUTOFS_DEV_IOCTL_EXPIRE_CMD** or the **AUTOFS_IOC_EXPIRE_MULTI** ioctl. This will work for both direct and indirect mounts. If it selects an object to expire, it will notify the daemon using the notification mechanism described below. This will block until the daemon acknowledges the expiry notification. This implies that the "*EXPIRE*" ioctl must be sent from a different thread than the one which handles notification.

While the ioctl is blocking, the entry is marked as "expiring" and *d_manage* will block until the daemon affirms that the unmount has completed (together with removing any directories that might have been necessary), or has been aborted.

### 3.5.6 Communicating with autofs: detecting the daemon

There are several forms of communication between the automount daemon and the filesystem. As we have already seen, the daemon can create and remove directories and symlinks using normal filesystem operations. autofs knows whether a process requesting some operation is the daemon or not based on its process-group id number (see getpgid(1)).

When an autofs filesystem is mounted the pgid of the mounting processes is recorded unless the "pgrp=" option is given, in which case that number is recorded instead. Any request arriving from a process in that process group is considered to come from the daemon. If the daemon ever has to be stopped and restarted a new pgid can be provided through an ioctl as will be described below.

### 3.5.7 Communicating with autofs: the event pipe

When an autofs filesystem is mounted, the 'write' end of a pipe must be passed using the 'fd=' mount option. autofs will write notification messages to this pipe for the daemon to respond to. For version 5, the format of the message is:

```
struct autofs_v5_packet {
        struct autofs_packet_hdr hdr;
        autofs_wqt_t wait_queue_token;
        __u32 dev;
        __u64 ino;
        __u32 uid;
        __u32 gid;
        __u32 pid;
        __u32 tgid;
        __u32 len;
        char name[NAME_MAX+1];
};
```

And the format of the header is:

```
struct autofs_packet_hdr {
        int proto_version;              /* Protocol version */
        int type;                       /* Type of packet */
};
```

where the type is one of

```
autofs_ptype_missing_indirect
autofs_ptype_expire_indirect
autofs_ptype_missing_direct
autofs_ptype_expire_direct
```

so messages can indicate that a name is missing (something tried to access it but it isn't there) or that it has been selected for expiry.

The pipe will be set to ”packet mode” (equivalent to passing *O_DIRECT*) to _pipe2(2)_ so that a read from the pipe will return at most one packet, and any unread portion of a packet will be discarded.

The *wait_queue_token* is a unique number which can identify a particular request to be acknowledged. When a message is sent over the pipe the affected dentry is marked as either ”active” or ”expiring” and other accesses to it block until the message is acknowledged using one of the ioctls below with the relevant *wait_queue_token*.

### 3.5.8 Communicating with autofs: root directory ioctls

The root directory of an autofs filesystem will respond to a number of ioctls. The process issuing the ioctl must have the CAP_SYS_ADMIN capability, or must be the automount daemon.

The available ioctl commands are:

- **AUTOFS_IOC_READY:**
  a notification has been handled. The argument to the ioctl command is the "wait_queue_token" number corresponding to the notification being acknowledged.
- **AUTOFS_IOC_FAIL:**
  similar to above, but indicates failure with the error code *ENOENT*.
- **AUTOFS_IOC_CATATONIC:**
  Causes the autofs to enter "catatonic" mode meaning that it stops sending notifications to the daemon. This mode is also entered if a write to the pipe fails.
- **AUTOFS_IOC_PROTOVER:**
  This returns the protocol version in use.
- **AUTOFS_IOC_PROTOSUBVER:**
  Returns the protocol sub-version which is really a version number for the implementation.
- **AUTOFS_IOC_SETTIMEOUT:**
  This passes a pointer to an unsigned long. The value is used to set the timeout for expiry, and the current timeout value is stored back through the pointer.
- **AUTOFS_IOC_ASKUMOUNT:**
  Returns, in the pointed-to *int*, 1 if the filesystem could be unmounted. This is only a hint as the situation could change at any instant. This call can be used to avoid a more expensive full unmount attempt.
- **AUTOFS_IOC_EXPIRE:**
  as described above, this asks if there is anything suitable to expire. A pointer to a packet:

  ```
  struct autofs_packet_expire_multi {
          struct autofs_packet_hdr hdr;
          autofs_wqt_t wait_queue_token;
          int len;
          char name[NAME_MAX+1];
  };
  ```

  is required. This is filled in with the name of something that can be unmounted or removed. If nothing can be expired, *errno* is set to *EAGAIN*. Even though a *wait_queue_token* is present in the structure, no "wait queue" is established and no acknowledgment is needed.
- **AUTOFS_IOC_EXPIRE_MULTI:**
  This is similar to **AUTOFS_IOC_EXPIRE** except that it causes notification to be sent to the daemon, and it blocks until the daemon acknowledges. The argument is an integer which can contain two different flags.

  **AUTOFS_EXP_IMMEDIATE** causes *last_used* time to be ignored and objects are expired if the are not in use.

**AUTOFS_EXP_FORCED** causes the in use status to be ignored and objects are expired ieven if they are in use. This assumes that the daemon has requested this because it is capable of performing the umount.

**AUTOFS_EXP_LEAVES** will select a leaf rather than a top-level name to expire. This is only safe when *maxproto* is 4.

### 3.5.9 Communicating with autofs: char-device ioctls

It is not always possible to open the root of an autofs filesystem, particularly a *direct* mounted filesystem. If the automount daemon is restarted there is no way for it to regain control of existing mounts using any of the above communication channels. To address this need there is a "miscellaneous" character device (major 10, minor 235) which can be used to communicate directly with the autofs filesystem. It requires CAP_SYS_ADMIN for access.

The 'ioctl's that can be used on this device are described in a separate document Miscellaneous Device control operations for the autofs kernel module, and are summarised briefly here. Each ioctl is passed a pointer to an *autofs_dev_ioctl* structure:

```
struct autofs_dev_ioctl {
        __u32 ver_major;
        __u32 ver_minor;
        __u32 size;             /* total size of data passed in
                                 * including this struct */
        __s32 ioctlfd;          /* automount command fd */

        /* Command parameters */
        union {
                struct args_protover            protover;
                struct args_protosubver         protosubver;
                struct args_openmount           openmount;
                struct args_ready               ready;
                struct args_fail                fail;
                struct args_setpipefd           setpipefd;
                struct args_timeout             timeout;
                struct args_requester           requester;
                struct args_expire              expire;
                struct args_askumount           askumount;
                struct args_ismountpoint        ismountpoint;
        };

        char path[];
};
```

For the **OPEN_MOUNT** and **IS_MOUNTPOINT** commands, the target filesystem is identified by the *path*. All other commands identify the filesystem by the *ioctlfd* which is a file descriptor open on the root, and which can be returned by **OPEN_MOUNT**.

The *ver_major* and *ver_minor* are in/out parameters which check that the requested version is supported, and report the maximum version that the kernel module can support.

Commands are:

- **AUTOFS_DEV_IOCTL_VERSION_CMD:**
  does nothing, except validate and set version numbers.
- **AUTOFS_DEV_IOCTL_OPENMOUNT_CMD:**
  return an open file descriptor on the root of an autofs filesystem. The filesystem is identified by name and device number, which is stored in *openmount.devid*. Device numbers for existing filesystems can be found in */proc/self/mountinfo*.
- **AUTOFS_DEV_IOCTL_CLOSEMOUNT_CMD:**
  same as *close(ioctlfd)*.
- **AUTOFS_DEV_IOCTL_SETPIPEFD_CMD:**
  if the filesystem is in catatonic mode, this can provide the write end of a new pipe in *setpipefd.pipefd* to re-establish communication with a daemon. The process group of the calling process is used to identify the daemon.
- **AUTOFS_DEV_IOCTL_REQUESTER_CMD:**
  *path* should be a name within the filesystem that has been auto-mounted on. On successful return, *requester.uid* and *requester.gid* will be the UID and GID of the process which triggered that mount.
- **AUTOFS_DEV_IOCTL_ISMOUNTPOINT_CMD:**
  Check if path is a mountpoint of a particular type - see separate documentation for details.
- **AUTOFS_DEV_IOCTL_PROTOVER_CMD**
- **AUTOFS_DEV_IOCTL_PROTOSUBVER_CMD**
- **AUTOFS_DEV_IOCTL_READY_CMD**
- **AUTOFS_DEV_IOCTL_FAIL_CMD**
- **AUTOFS_DEV_IOCTL_CATATONIC_CMD**
- **AUTOFS_DEV_IOCTL_TIMEOUT_CMD**
- **AUTOFS_DEV_IOCTL_EXPIRE_CMD**
- **AUTOFS_DEV_IOCTL_ASKUMOUNT_CMD**

These all have the same function as the similarly named **AUTOFS_IOC** ioctls, except that **FAIL** can be given an explicit error number in *fail.status* instead of assuming *ENOENT*, and this **EXPIRE** command corresponds to **AUTOFS_IOC_EXPIRE_MULTI**.

### 3.5.10 Catatonic mode

As mentioned, an autofs mount can enter ”catatonic” mode. This happens if a write to the notification pipe fails, or if it is explicitly requested by an *ioctl*.

When entering catatonic mode, the pipe is closed and any pending notifications are acknowledged with the error *ENOENT*.

Once in catatonic mode attempts to access non-existing names will result in *ENOENT* while attempts to access existing directories will be treated in the same way as if they came from the daemon, so mount traps will not fire.

When the filesystem is mounted a _uid_ and _gid_ can be given which set the ownership of directories and symbolic links. When the filesystem is in catatonic mode, any process with a

matching UID can create directories or symlinks in the root directory, but not in other directories.

Catatonic mode can only be left via the **AUTOFS_DEV_IOCTL_OPENMOUNT_CMD** ioctl on the */dev/autofs*.

### 3.5.11 The ”ignore” mount option

The ”ignore” mount option can be used to provide a generic indicator to applications that the mount entry should be ignored when displaying mount information.

In other OSes that provide autofs and that provide a mount list to user space based on the kernel mount list a no-op mount option (”ignore” is the one use on the most common OSes) is allowed so that autofs file system users can optionally use it.

This is intended to be used by user space programs to exclude autofs mounts from consideration when reading the mounts list.

### 3.5.12 autofs, name spaces, and shared mounts

With bind mounts and name spaces it is possible for an autofs filesystem to appear at multiple places in one or more filesystem name spaces. For this to work sensibly, the autofs filesystem should always be mounted ”shared”. e.g.

```
mount --make-shared /autofs/mount/point
```

The automount daemon is only able to manage a single mount location for an autofs filesystem and if mounts on that are not 'shared', other locations will not behave as expected. In particular access to those other locations will likely result in the *ELOOP* error

```
Too many levels of symbolic links
```

## 3.6 Miscellaneous Device control operations for the autofs kernel module

### 3.6.1 The problem

There is a problem with active restarts in autofs (that is to say restarting autofs when there are busy mounts).

During normal operation autofs uses a file descriptor opened on the directory that is being managed in order to be able to issue control operations. Using a file descriptor gives ioctl operations access to autofs specific information stored in the super block. The operations are things such as setting an autofs mount catatonic, setting the expire timeout and requesting expire checks. As is explained below, certain types of autofs triggered mounts can end up covering an autofs mount itself which prevents us being able to use open(2) to obtain a file descriptor for these operations if we don't already have one open.

Currently autofs uses ”umount -l” (lazy umount) to clear active mounts at restart. While using lazy umount works for most cases, anything that needs to walk back up the mount tree to

construct a path, such as getcwd(2) and the proc file system /proc/<pid>/cwd, no longer works because the point from which the path is constructed has been detached from the mount tree.

The actual problem with autofs is that it can't reconnect to existing mounts. Immediately one thinks of just adding the ability to remount autofs file systems would solve it, but alas, that can't work. This is because autofs direct mounts and the implementation of ”on demand mount and expire” of nested mount trees have the file system mounted directly on top of the mount trigger directory dentry.

For example, there are two types of automount maps, direct (in the kernel module source you will see a third type called an offset, which is just a direct mount in disguise) and indirect.

Here is a master map with direct and indirect map entries:

```
/-      /etc/auto.direct
/test   /etc/auto.indirect
```

and the corresponding map files:

```
/etc/auto.direct:

/automount/dparse/g6  budgie:/autofs/export1
/automount/dparse/g1  shark:/autofs/export1
and so on.
```

/etc/auto.indirect:

```
g1    shark:/autofs/export1
g6    budgie:/autofs/export1
and so on.
```

For the above indirect map an autofs file system is mounted on /test and mounts are triggered for each sub-directory key by the inode lookup operation. So we see a mount of shark:/autofs/export1 on /test/g1, for example.

The way that direct mounts are handled is by making an autofs mount on each full path, such as /automount/dparse/g1, and using it as a mount trigger. So when we walk on the path we mount shark:/autofs/export1 ”on top of this mount point”. Since these are always directories we can use the follow_link inode operation to trigger the mount.

But, each entry in direct and indirect maps can have offsets (making them multi-mount map entries).

For example, an indirect mount map entry could also be:

```
g1  \
/         shark:/autofs/export5/testing/test \
/s1       shark:/autofs/export/testing/test/s1 \
/s2       shark:/autofs/export5/testing/test/s2 \
/s1/ss1   shark:/autofs/export1 \
/s2/ss2   shark:/autofs/export2
```

and a similarly a direct mount map entry could also be:

```
/automount/dparse/g1 \
    /       shark:/autofs/export5/testing/test \
    /s1     shark:/autofs/export/testing/test/s1 \
    /s2     shark:/autofs/export5/testing/test/s2 \
    /s1/ss1 shark:/autofs/export2 \
    /s2/ss2 shark:/autofs/export2
```

One of the issues with version 4 of autofs was that, when mounting an entry with a large number of offsets, possibly with nesting, we needed to mount and umount all of the offsets as a single unit. Not really a problem, except for people with a large number of offsets in map entries. This mechanism is used for the well known ”hosts” map and we have seen cases (in 2.4) where the available number of mounts are exhausted or where the number of privileged ports available is exhausted.

In version 5 we mount only as we go down the tree of offsets and similarly for expiring them which resolves the above problem. There is somewhat more detail to the implementation but it isn't needed for the sake of the problem explanation. The one important detail is that these offsets are implemented using the same mechanism as the direct mounts above and so the mount points can be covered by a mount.

The current autofs implementation uses an ioctl file descriptor opened on the mount point for control operations. The references held by the descriptor are accounted for in checks made to determine if a mount is in use and is also used to access autofs file system information held in the mount super block. So the use of a file handle needs to be retained.

### 3.6.2 The Solution

To be able to restart autofs leaving existing direct, indirect and offset mounts in place we need to be able to obtain a file handle for these potentially covered autofs mount points. Rather than just implement an isolated operation it was decided to re-implement the existing ioctl interface and add new operations to provide this functionality.

In addition, to be able to reconstruct a mount tree that has busy mounts, the uid and gid of the last user that triggered the mount needs to be available because these can be used as macro substitution variables in autofs maps. They are recorded at mount request time and an operation has been added to retrieve them.

Since we're re-implementing the control interface, a couple of other problems with the existing interface have been addressed. First, when a mount or expire operation completes a status is returned to the kernel by either a ”send ready” or a ”send fail” operation. The ”send fail” operation of the ioctl interface could only ever send ENOENT so the re-implementation allows user space to send an actual status. Another expensive operation in user space, for those using very large maps, is discovering if a mount is present. Usually this involves scanning /proc/mounts and since it needs to be done quite often it can introduce significant overhead when there are many entries in the mount table. An operation to lookup the mount status of a mount point dentry (covered or not) has also been added.

Current kernel development policy recommends avoiding the use of the ioctl mechanism in favor of systems such as Netlink. An implementation using this system was attempted to evaluate its suitability and it was found to be inadequate, in this case. The Generic Netlink system was used for this as raw Netlink would lead to a significant increase in complexity. There's no question that the Generic Netlink system is an elegant solution for common case ioctl functions but it's not a complete replacement probably because its primary purpose in life is to be a message bus

implementation rather than specifically an ioctl replacement. While it would be possible to work around this there is one concern that lead to the decision to not use it. This is that the autofs expire in the daemon has become far to complex because umount candidates are enumerated, almost for no other reason than to ”count” the number of times to call the expire ioctl. This involves scanning the mount table which has proved to be a big overhead for users with large maps. The best way to improve this is try and get back to the way the expire was done long ago. That is, when an expire request is issued for a mount (file handle) we should continually call back to the daemon until we can't umount any more mounts, then return the appropriate status to the daemon. At the moment we just expire one mount at a time. A Generic Netlink implementation would exclude this possibility for future development due to the requirements of the message bus architecture.

### 3.6.3 autofs Miscellaneous Device mount control interface

The control interface is opening a device node, typically /dev/autofs.

All the ioctls use a common structure to pass the needed parameter information and return operation results:

```
struct autofs_dev_ioctl {
        __u32 ver_major;
        __u32 ver_minor;
        __u32 size;             /* total size of data passed in
                                * including this struct */
        __s32 ioctlfd;          /* automount command fd */

        /* Command parameters */
        union {
                struct args_protover                protover;
                struct args_protosubver             protosubver;
                struct args_openmount               openmount;
                struct args_ready           ready;
                struct args_fail            fail;
                struct args_setpipefd               setpipefd;
                struct args_timeout         timeout;
                struct args_requester               requester;
                struct args_expire          expire;
                struct args_askumount               askumount;
                struct args_ismountpoint    ismountpoint;
        };

        char path[];
};
```

The ioctlfd field is a mount point file descriptor of an autofs mount point. It is returned by the open call and is used by all calls except the check for whether a given path is a mount point, where it may optionally be used to check a specific mount corresponding to a given mount point file descriptor, and when requesting the uid and gid of the last successful mount on a directory within the autofs file system.

The union is used to communicate parameters and results of calls made as described below.

The path field is used to pass a path where it is needed and the size field is used account for the increased structure length when translating the structure sent from user space.

This structure can be initialized before setting specific fields by using the void function call init_autofs_dev_ioctl(`struct autofs_dev_ioctl *`).

All of the ioctls perform a copy of this structure from user space to kernel space and return -EINVAL if the size parameter is smaller than the structure size itself, -ENOMEM if the kernel memory allocation fails or -EFAULT if the copy itself fails. Other checks include a version check of the compiled in user space version against the module version and a mismatch results in a -EINVAL return. If the size field is greater than the structure size then a path is assumed to be present and is checked to ensure it begins with a "/" and is NULL terminated, otherwise -EINVAL is returned. Following these checks, for all ioctl commands except AUTOFS_DEV_IOCTL_VERSION_CMD, AUTOFS_DEV_IOCTL_OPENMOUNT_CMD and AUTOFS_DEV_IOCTL_CLOSEMOUNT_CMD the ioctlfd is validated and if it is not a valid descriptor or doesn't correspond to an autofs mount point an error of -EBADF, -ENOTTY or -EINVAL (not an autofs descriptor) is returned.

### 3.6.4 The ioctls

An example of an implementation which uses this interface can be seen in autofs version 5.0.4 and later in file lib/dev-ioctl-lib.c of the distribution tar available for download from kernel.org in directory /pub/linux/daemons/autofs/v5.

The device node ioctl operations implemented by this interface are:

#### AUTOFS_DEV_IOCTL_VERSION

Get the major and minor version of the autofs device ioctl kernel module implementation. It requires an initialized struct autofs_dev_ioctl as an input parameter and sets the version information in the passed in structure. It returns 0 on success or the error -EINVAL if a version mismatch is detected.

### AUTOFS_DEV_IOCTL_PROTOVER_CMD and AUTOFS_DEV_IOCTL_PROTOSUBVER_CMD

Get the major and minor version of the autofs protocol version understood by loaded module. This call requires an initialized struct autofs_dev_ioctl with the ioctlfd field set to a valid autofs mount point descriptor and sets the requested version number in version field of struct args_protover or sub_version field of struct args_protosubver. These commands return 0 on success or one of the negative error codes if validation fails.

### AUTOFS_DEV_IOCTL_OPENMOUNT and AUTOFS_DEV_IOCTL_CLOSEMOUNT

Obtain and release a file descriptor for an autofs managed mount point path. The open call requires an initialized struct autofs_dev_ioctl with the path field set and the size field adjusted appropriately as well as the devid field of struct args_openmount set to the device number of the autofs mount. The device number can be obtained from the mount options shown in /proc/mounts. The close call requires an initialized struct autofs_dev_ioct with the ioctlfd field set to the descriptor obtained from the open call. The release of the file descriptor can also be done with close(2) so any open descriptors will also be closed at process exit. The close call is included in the implemented operations largely for completeness and to provide for a consistent user space implementation.

### AUTOFS_DEV_IOCTL_READY_CMD and AUTOFS_DEV_IOCTL_FAIL_CMD

Return mount and expire result status from user space to the kernel. Both of these calls require an initialized struct autofs_dev_ioctl with the ioctlfd field set to the descriptor obtained from the open call and the token field of struct args_ready or struct args_fail set to the wait queue token number, received by user space in the foregoing mount or expire request. The status field of struct args_fail is set to the errno of the operation. It is set to 0 on success.

### AUTOFS_DEV_IOCTL_SETPIPEFD_CMD

Set the pipe file descriptor used for kernel communication to the daemon. Normally this is set at mount time using an option but when reconnecting to a existing mount we need to use this to tell the autofs mount about the new kernel pipe descriptor. In order to protect mounts against incorrectly setting the pipe descriptor we also require that the autofs mount be catatonic (see next call).

The call requires an initialized struct autofs_dev_ioctl with the ioctlfd field set to the descriptor obtained from the open call and the pipefd field of struct args_setpipefd set to descriptor of the pipe. On success the call also sets the process group id used to identify the controlling process (eg. the owning automount(8) daemon) to the process group of the caller.

### AUTOFS_DEV_IOCTL_CATATONIC_CMD

Make the autofs mount point catatonic. The autofs mount will no longer issue mount requests, the kernel communication pipe descriptor is released and any remaining waits in the queue released.

The call requires an initialized struct autofs_dev_ioctl with the ioctlfd field set to the descriptor obtained from the open call.

### AUTOFS_DEV_IOCTL_TIMEOUT_CMD

Set the expire timeout for mounts within an autofs mount point.

The call requires an initialized struct autofs_dev_ioctl with the ioctlfd field set to the descriptor obtained from the open call.

### AUTOFS_DEV_IOCTL_REQUESTER_CMD

Return the uid and gid of the last process to successfully trigger a the mount on the given path dentry.

The call requires an initialized struct autofs_dev_ioctl with the path field set to the mount point in question and the size field adjusted appropriately. Upon return the uid field of struct args_requester contains the uid and gid field the gid.

When reconstructing an autofs mount tree with active mounts we need to re-connect to mounts that may have used the original process uid and gid (or string variations of them) for mount lookups within the map entry. This call provides the ability to obtain this uid and gid so they may be used by user space for the mount map lookups.

### AUTOFS_DEV_IOCTL_EXPIRE_CMD

Issue an expire request to the kernel for an autofs mount. Typically this ioctl is called until no further expire candidates are found.

The call requires an initialized struct autofs_dev_ioctl with the ioctlfd field set to the descriptor obtained from the open call. In addition an immediate expire that's independent of the mount timeout, and a forced expire that's independent of whether the mount is busy, can be requested by setting the how field of struct args_expire to AUTOFS_EXP_IMMEDIATE or AUTOFS_EXP_FORCED, respectively . If no expire candidates can be found the ioctl returns -1 with errno set to EAGAIN.

This call causes the kernel module to check the mount corresponding to the given ioctlfd for mounts that can be expired, issues an expire request back to the daemon and waits for completion.

#### AUTOFS_DEV_IOCTL_ASKUMOUNT_CMD

Checks if an autofs mount point is in use.

The call requires an initialized struct autofs_dev_ioctl with the ioctlfd field set to the descriptor obtained from the open call and it returns the result in the may_umount field of struct args_askumount, 1 for busy and 0 otherwise.

#### AUTOFS_DEV_IOCTL_ISMOUNTPOINT_CMD

Check if the given path is a mountpoint.

The call requires an initialized struct autofs_dev_ioctl. There are two possible variations. Both use the path field set to the path of the mount point to check and the size field adjusted appropriately. One uses the ioctlfd field to identify a specific mount point to check while the other variation uses the path and optionally in.type field of struct args_ismountpoint set to an autofs mount type. The call returns 1 if this is a mount point and sets out.devid field to the device number of the mount and out.magic field to the relevant super block magic number (described below) or 0 if it isn't a mountpoint. In both cases the device number (as returned by new_encode_dev()) is returned in out.devid field.

If supplied with a file descriptor we're looking for a specific mount, not necessarily at the top of the mounted stack. In this case the path the descriptor corresponds to is considered a mountpoint if it is itself a mountpoint or contains a mount, such as a multi-mount without a root mount. In this case we return 1 if the descriptor corresponds to a mount point and also returns the super magic of the covering mount if there is one or 0 if it isn't a mountpoint.

If a path is supplied (and the ioctlfd field is set to -1) then the path is looked up and is checked to see if it is the root of a mount. If a type is also given we are looking for a particular autofs mount and if a match isn't found a fail is returned. If the located path is the root of a mount 1 is returned along with the super magic of the mount or 0 otherwise.

## 3.7 BeOS filesystem for Linux

Document last updated: Dec 6, 2001

### 3.7.1 Warning

Make sure you understand that this is alpha software. This means that the implementation is neither complete nor well-tested.

I DISCLAIM ALL RESPONSIBILITY FOR ANY POSSIBLE BAD EFFECTS OF THIS CODE!

### 3.7.2 License

This software is covered by the GNU General Public License. See the file COPYING for the complete text of the license. Or the GNU website: <http://www.gnu.org/licenses/licenses.html>

### 3.7.3 Author

The largest part of the code written by Will Dyson <will_dyson@pobox.com> He has been working on the code since Aug 13, 2001. See the changelog for details.

Original Author: Makoto Kato <m_kato@ga2.so-net.ne.jp>

His original code can still be found at: <http://hp.vector.co.jp/authors/VA008030/bfs/>

Does anyone know of a more current email address for Makoto? He doesn't respond to the address given above...

This filesystem doesn't have a maintainer.

### 3.7.4 What is this Driver?

This module implements the native filesystem of BeOS http://www.beincorporated.com/ for the linux 2.4.1 and later kernels. Currently it is a read-only implementation.

### 3.7.5 Which is it, BFS or BEFS?

Be, Inc said, "BeOS Filesystem is officially called BFS, not BeFS". But Unixware Boot Filesystem is called bfs, too. And they are already in the kernel. Because of this naming conflict, on Linux the BeOS filesystem is called befs.

### 3.7.6 How to Install

step 1. Install the BeFS patch into the source code tree of linux.

Apply the patchfile to your kernel source tree. Assuming that your kernel source is in /foo/bar/linux and the patchfile is called patch-befs-xxx, you would do the following:

> cd /foo/bar/linux patch -p1 < /path/to/patch-befs-xxx

if the patching step fails (i.e. there are rejected hunks), you can try to figure it out yourself (it shouldn't be hard), or mail the maintainer (Will Dyson <will_dyson@pobox.com>) for help.

step 2. Configuration & make kernel

The linux kernel has many compile-time options. Most of them are beyond the scope of this document. I suggest the Kernel-HOWTO document as a good general reference on this topic. http://www.linuxdocs.org/HOWTOs/Kernel-HOWTO-4.html

However, to use the BeFS module, you must enable it at configure time:

```
cd /foo/bar/linux
make menuconfig (or xconfig)
```

The BeFS module is not a standard part of the linux kernel, so you must first enable support for experimental code under the ”Code maturity level” menu.

Then, under the ”Filesystems” menu will be an option called ”BeFS filesystem (experimental)”, or something like that. Enable that option (it is fine to make it a module).

Save your kernel configuration and then build your kernel.

step 3. Install

See the kernel howto <http://www.linux.com/howto/Kernel-HOWTO.html> for instructions on this critical step.

### 3.7.7 Using BFS

To use the BeOS filesystem, use filesystem type 'befs'.

ex:

```
mount -t befs /dev/fd0 /beos
```

### 3.7.8 Mount Options

| | |
|---|---|
| uid=nnn | All files in the partition will be owned by user id nnn. |
| gid=nnn | All files in the partition will be in group nnn. |
| iocharset=xxx | Use xxx as the name of the NLS translation table. |
| debug | The driver will output debugging information to the syslog. |

### 3.7.9 How to Get Latest Version

The latest version is currently available at: <http://befs-driver.sourceforge.net/>

### 3.7.10 Any Known Bugs?

As of Jan 20, 2002:

None

### 3.7.11 Special Thanks

Dominic Giampalo ... Writing ”Practical file system design with Be filesystem”

Hiroyuki Yamada ... Testing LinuxPPC.

## 3.8 BFS Filesystem for Linux

The BFS filesystem is used by SCO UnixWare OS for the /stand slice, which usually contains the kernel image and a few other files required for the boot process.

In order to access /stand partition under Linux you obviously need to know the partition number and the kernel must support UnixWare disk slices (CONFIG_UNIXWARE_DISKLABEL config option). However BFS support does not depend on having UnixWare disklabel support because one can also mount BFS filesystem via loopback:

```
# losetup /dev/loop0 stand.img
# mount -t bfs /dev/loop0 /mnt/stand
```

where stand.img is a file containing the image of BFS filesystem. When you have finished using it and umounted you need to also deallocate /dev/loop0 device by:

```
# losetup -d /dev/loop0
```

You can simplify mounting by just typing:

```
# mount -t bfs -o loop stand.img /mnt/stand
```

this will allocate the first available loopback device (and load loop.o kernel module if necessary) automatically. If the loopback driver is not loaded automatically, make sure that you have compiled the module and that modprobe is functioning. Beware that umount will not deallocate /dev/loopN device if /etc/mtab file on your system is a symbolic link to /proc/mounts. You will need to do it manually using ”-d” switch of losetup(8). Read losetup(8) manpage for more info.

To create the BFS image under UnixWare you need to find out first which slice contains it. The command prtvtoc(1M) is your friend:

```
# prtvtoc /dev/rdsk/c0b0t0d0s0
```

(assuming your root disk is on target=0, lun=0, bus=0, controller=0). Then you look for the slice with tag ”STAND”, which is usually slice 10. With this information you can use dd(1) to create the BFS image:

```
# umount /stand
# dd if=/dev/rdsk/c0b0t0d0sa of=stand.img bs=512
```

Just in case, you can verify that you have done the right thing by checking the magic number:

```
# od -Ad -tx4 stand.img | more
```

The first 4 bytes should be 0x1badface.

If you have any patches, questions or suggestions regarding this BFS implementation please contact the author:

Tigran Aivazian <aivazian.tigran@gmail.com>

## 3.9 BTRFS

Btrfs is a copy on write filesystem for Linux aimed at implementing advanced features while focusing on fault tolerance, repair and easy administration. Jointly developed by several companies, licensed under the GPL and open for contribution from anyone.

The main Btrfs features include:

- Extent based file storage (2^64 max file size)
- Space efficient packing of small files
- Space efficient indexed directories
- Dynamic inode allocation
- Writable snapshots
- Subvolumes (separate internal filesystem roots)
- Object level mirroring and striping
- Checksums on data and metadata (multiple algorithms available)
- Compression (multiple algorithms available)
- Reflink, deduplication
- Scrub (on-line checksum verification)
- Hierarchical quota groups (subvolume and snapshot support)
- Integrated multiple device support, with several raid algorithms
- Offline filesystem check
- Efficient incremental backup and FS mirroring (send/receive)
- Trim/discard
- Online filesystem defragmentation
- Swapfile support
- Zoned mode
- Read/write metadata verification
- Online resize (shrink, grow)

For more information please refer to the documentation site or wiki

https://btrfs.readthedocs.io

that maintains information about administration tasks, frequently asked questions, use cases, mount options, comprehensible changelogs, features, manual pages, source code repositories, contacts etc.

## 3.10 Ceph Distributed File System

Ceph is a distributed network file system designed to provide good performance, reliability, and scalability.

Basic features include:

- POSIX semantics
- Seamless scaling from 1 to many thousands of nodes
- High availability and reliability. No single point of failure.
- N-way replication of data across storage nodes
- Fast recovery from node failures
- Automatic rebalancing of data on node addition/removal
- Easy deployment: most FS components are userspace daemons

Also,

- Flexible snapshots (on any directory)
- Recursive accounting (nested files, directories, bytes)

In contrast to cluster filesystems like GFS, OCFS2, and GPFS that rely on symmetric access by all clients to shared block devices, Ceph separates data and metadata management into independent server clusters, similar to Lustre. Unlike Lustre, however, metadata and storage nodes run entirely as user space daemons. File data is striped across storage nodes in large chunks to distribute workload and facilitate high throughputs. When storage nodes fail, data is re-replicated in a distributed fashion by the storage nodes themselves (with some minimal coordination from a cluster monitor), making the system extremely efficient and scalable.

Metadata servers effectively form a large, consistent, distributed in-memory cache above the file namespace that is extremely scalable, dynamically redistributes metadata in response to workload changes, and can tolerate arbitrary (well, non-Byzantine) node failures. The metadata server takes a somewhat unconventional approach to metadata storage to significantly improve performance for common workloads. In particular, inodes with only a single link are embedded in directories, allowing entire directories of dentries and inodes to be loaded into its cache with a single I/O operation. The contents of extremely large directories can be fragmented and managed by independent metadata servers, allowing scalable concurrent access.

The system offers automatic data rebalancing/migration when scaling from a small cluster of just a few nodes to many hundreds, without requiring an administrator carve the data set into static volumes or go through the tedious process of migrating data between servers. When the file system approaches full, new nodes can be easily added and things will ”just work.”

Ceph includes flexible snapshot mechanism that allows a user to create a snapshot on any subdirectory (and its nested contents) in the system. Snapshot creation and deletion are as simple as 'mkdir .snap/foo' and 'rmdir .snap/foo'.

Snapshot names have two limitations:

- They can not start with an underscore ('_'), as these names are reserved for internal usage by the MDS.

- They can not exceed 240 characters in size. This is because the MDS makes use of long snapshot names internally, which follow the format: _<*SNAPSHOT-NAME*>_<*INODE-NUMBER*>. Since filenames in general can't have more than 255 characters, and <*node-id*> takes 13 characters, the long snapshot names can take as much as 255 - 1 - 1 - 13 = 240.

Ceph also provides some recursive accounting on directories for nested files and bytes. That is, a 'getfattr -d foo' on any directory in the system will reveal the total number of nested regular files and subdirectories, and a summation of all nested file sizes. This makes the identification of large disk space consumers relatively quick, as no 'du' or similar recursive scan of the file system is required.

Finally, Ceph also allows quotas to be set on any directory in the system. The quota can restrict the number of bytes or the number of files stored beneath that point in the directory hierarchy. Quotas can be set using extended attributes 'ceph.quota.max_files' and 'ceph.quota.max_bytes', eg:

```
setfattr -n ceph.quota.max_bytes -v 100000000 /some/dir
getfattr -n ceph.quota.max_bytes /some/dir
```

A limitation of the current quotas implementation is that it relies on the cooperation of the client mounting the file system to stop writers when a limit is reached. A modified or adversarial client cannot be prevented from writing as much data as it needs.

### 3.10.1 Mount Syntax

The basic mount syntax is:

```
# mount -t ceph user@fsid.fs_name=/[subdir] mnt -o mon_addr=monip1[:port][/
 ↪monip2[:port]]
```

You only need to specify a single monitor, as the client will get the full list when it connects. (However, if the monitor you specify happens to be down, the mount won't succeed.) The port can be left off if the monitor is using the default. So if the monitor is at 1.2.3.4:

```
# mount -t ceph cephuser@07fe3187-00d9-42a3-814b-72a4d5e7d5be.cephfs=/ /mnt/
 ↪ceph -o mon_addr=1.2.3.4
```

is sufficient. If /sbin/mount.ceph is installed, a hostname can be used instead of an IP address and the cluster FSID can be left out (as the mount helper will fill it in by reading the ceph configuration file):

```
# mount -t ceph cephuser@cephfs=/ /mnt/ceph -o mon_addr=mon-addr
```

Multiple monitor addresses can be passed by separating each address with a slash (/):

```
# mount -t ceph cephuser@cephfs=/ /mnt/ceph -o mon_addr=192.168.1.100/192.168.
 ↪1.101
```

When using the mount helper, monitor address can be read from ceph configuration file if available. Note that, the cluster FSID (passed as part of the device string) is validated by checking it with the FSID reported by the monitor.

### 3.10.2 Mount Options

**mon_addr=ip_address[:port][/ip_address[:port]]**
Monitor address to the cluster. This is used to bootstrap the connection to the cluster. Once connection is established, the monitor addresses in the monitor map are followed.

**fsid=cluster-id**
FSID of the cluster (from *ceph fsid* command).

**ip=A.B.C.D[:N]**
Specify the IP and/or port the client should bind to locally. There is normally not much reason to do this. If the IP is not specified, the client's IP address is determined by looking at the address its connection to the monitor originates from.

**wsize=X**
Specify the maximum write size in bytes. Default: 64 MB.

**rsize=X**
Specify the maximum read size in bytes. Default: 64 MB.

**rasize=X**
Specify the maximum readahead size in bytes. Default: 8 MB.

**mount_timeout=X**
Specify the timeout value for mount (in seconds), in the case of a non-responsive Ceph file system. The default is 60 seconds.

**caps_max=X**
Specify the maximum number of caps to hold. Unused caps are released when number of caps exceeds the limit. The default is 0 (no limit)

**rbytes**
When stat() is called on a directory, set st_size to 'rbytes', the summation of file sizes over all files nested beneath that directory. This is the default.

**norbytes**
When stat() is called on a directory, set st_size to the number of entries in that directory.

**nocrc**
Disable CRC32C calculation for data writes. If set, the storage node must rely on TCP's error correction to detect data corruption in the data payload.

**dcache**
Use the dcache contents to perform negative lookups and readdir when the client has the entire directory contents in its cache. (This does not change correctness; the client uses cached metadata only when a lease or capability ensures it is valid.)

**nodcache**
Do not use the dcache as above. This avoids a significant amount of complex code, sacrificing performance without affecting correctness, and is useful for tracking down bugs.

**noasyncreaddir**
Do not use the dcache as above for readdir.

**noquotadf**
Report overall filesystem usage in statfs instead of using the root directory quota.

**nocopyfrom**
Don't use the RADOS 'copy-from' operation to perform remote object copies. Currently, it's only used in copy_file_range, which will revert to the default VFS implementation if this option is used.

**recover_session=<no|clean>**
Set auto reconnect mode in the case where the client is blocklisted. The available modes are ”no” and ”clean”. The default is ”no”.

- no: never attempt to reconnect when client detects that it has been blocklisted. Operations will generally fail after being blocklisted.
- clean: client reconnects to the ceph cluster automatically when it detects that it has been blocklisted. During reconnect, client drops dirty data/metadata, invalidates page caches and writable file handles. After reconnect, file locks become stale because the MDS loses track of them. If an inode contains any stale file locks, read/write on the inode is not allowed until applications release all stale file locks.

### 3.10.3 More Information

**For more information on Ceph, see the home page at**
https://ceph.com/

**The Linux kernel client source tree is available at**

- https://github.com/ceph/ceph-client.git

**and the source for the full system is at**
https://github.com/ceph/ceph.git

## 3.11 Coda Kernel-Venus Interface

---

**Note:** This is one of the technical documents describing a component of Coda -- this document describes the client kernel-Venus interface.

---

For more information:

http://www.coda.cs.cmu.edu

For user level software needed to run Coda:

ftp://ftp.coda.cs.cmu.edu

To run Coda you need to get a user level cache manager for the client, named Venus, as well as tools to manipulate ACLs, to log in, etc. The client needs to have the Coda filesystem selected in the kernel configuration.

The server needs a user level server and at present does not depend on kernel support.

The Venus kernel interface

Peter J. Braam

v1.0, Nov 9, 1997

This document describes the communication between Venus and kernel level filesystem code needed for the operation of the Coda file sys- tem. This document version is meant to describe the current interface (version 1.0) as well as improvements we envisage.

### 3.11.1 1. Introduction

A key component in the Coda Distributed File System is the cache manager, Venus.

When processes on a Coda enabled system access files in the Coda filesystem, requests are directed at the filesystem layer in the operating system. The operating system will communicate with Venus to service the request for the process. Venus manages a persistent client cache and makes remote procedure calls to Coda file servers and related servers (such as authentication servers) to service these requests it receives from the operating system. When Venus has serviced a request it replies to the operating system with appropriate return codes, and other data related to the request. Optionally the kernel support for Coda may maintain a minicache of recently processed requests to limit the number of interactions with Venus. Venus possesses the facility to inform the kernel when elements from its minicache are no longer valid.

This document describes precisely this communication between the kernel and Venus. The definitions of so called upcalls and downcalls will be given with the format of the data they handle. We shall also describe the semantic invariants resulting from the calls.

Historically Coda was implemented in a BSD file system in Mach 2.6. The interface between the kernel and Venus is very similar to the BSD VFS interface. Similar functionality is provided, and the format of the parameters and returned data is very similar to the BSD VFS. This leads to an almost natural environment for implementing a kernel-level filesystem driver for Coda in a BSD system. However, other operating systems such as Linux and Windows 95 and NT have virtual filesystem with different interfaces.

To implement Coda on these systems some reverse engineering of the Venus/Kernel protocol is necessary. Also it came to light that other systems could profit significantly from certain small optimizations and modifications to the protocol. To facilitate this work as well as to make future ports easier, communication between Venus and the kernel should be documented in great detail. This is the aim of this document.

### 3.11.2 2. Servicing Coda filesystem calls

The service of a request for a Coda file system service originates in a process P which accessing a Coda file. It makes a system call which traps to the OS kernel. Examples of such calls trapping to the kernel are `read`, `write`, `open`, `close`, `create`, `mkdir`, `rmdir`, `chmod` in a Unix ontext. Similar calls exist in the Win32 environment, and are named `CreateFile`.

Generally the operating system handles the request in a virtual filesystem (VFS) layer, which is named I/O Manager in NT and IFS manager in Windows 95. The VFS is responsible for partial processing of the request and for locating the specific filesystem(s) which will service parts of the request. Usually the information in the path assists in locating the correct FS drivers. Sometimes after extensive pre-processing, the VFS starts invoking exported routines in the FS driver. This is the point where the FS specific processing of the request starts, and here the Coda specific kernel code comes into play.

The FS layer for Coda must expose and implement several interfaces. First and foremost the VFS must be able to make all necessary calls to the Coda FS layer, so the Coda FS driver must expose the VFS interface as applicable in the operating system. These differ very significantly among operating systems, but share features such as facilities to read/write and create and remove objects. The Coda FS layer services such VFS requests by invoking one or more well defined services offered by the cache manager Venus. When the replies from Venus have come back to the FS driver, servicing of the VFS call continues and finishes with a reply to the kernel's VFS. Finally the VFS layer returns to the process.

As a result of this design a basic interface exposed by the FS driver must allow Venus to manage message traffic. In particular Venus must be able to retrieve and place messages and to be notified of the arrival of a new message. The notification must be through a mechanism which does not block Venus since Venus must attend to other tasks even when no messages are waiting or being processed.

**Interfaces of the Coda FS Driver**

Furthermore the FS layer provides for a special path of communication between a user process and Venus, called the pioctl interface. The pioctl interface is used for Coda specific services, such as requesting detailed information about the persistent cache managed by Venus. Here the involvement of the kernel is minimal. It identifies the calling process and passes the information on to Venus. When Venus replies the response is passed back to the caller in unmodified form.

Finally Venus allows the kernel FS driver to cache the results from certain services. This is done to avoid excessive context switches and results in an efficient system. However, Venus may acquire information, for example from the network which implies that cached information must be flushed or replaced. Venus then makes a downcall to the Coda FS layer to request flushes or updates in the cache. The kernel FS driver handles such requests synchronously.

Among these interfaces the VFS interface and the facility to place, receive and be notified of messages are platform specific. We will not go into the calls exported to the VFS layer but we will state the requirements of the message exchange mechanism.

### 3.11.3 3. The message layer

At the lowest level the communication between Venus and the FS driver proceeds through messages. The synchronization between processes requesting Coda file service and Venus relies on blocking and waking up processes. The Coda FS driver processes VFS- and pioctl-requests on behalf of a process P, creates messages for Venus, awaits replies and finally returns to the caller. The implementation of the exchange of messages is platform specific, but the semantics have (so far) appeared to be generally applicable. Data buffers are created by the FS Driver in kernel memory on behalf of P and copied to user memory in Venus.

The FS Driver while servicing P makes upcalls to Venus. Such an upcall is dispatched to Venus by creating a message structure. The structure contains the identification of P, the message sequence number, the size of the request and a pointer to the data in kernel memory for the request. Since the data buffer is re-used to hold the reply from Venus, there is a field for the size of the reply. A flags field is used in the message to precisely record the status of the message. Additional platform dependent structures involve pointers to determine the position of the message on queues and pointers to synchronization objects. In the upcall routine the message structure is filled in, flags are set to 0, and it is placed on the *pending* queue. The routine calling upcall is responsible for allocating the data buffer; its structure will be described in the next section.

A facility must exist to notify Venus that the message has been created, and implemented using available synchronization objects in the OS. This notification is done in the upcall context of the process P. When the message is on the pending queue, process P cannot proceed in upcall. The (kernel mode) processing of P in the filesystem request routine must be suspended until Venus has replied. Therefore the calling thread in P is blocked in upcall. A pointer in the message structure will locate the synchronization object on which P is sleeping.

Venus detects the notification that a message has arrived, and the FS driver allow Venus to retrieve the message with a getmsg_from_kernel call. This action finishes in the kernel by putting the message on the queue of processing messages and setting flags to READ. Venus is passed the contents of the data buffer. The getmsg_from_kernel call now returns and Venus processes the request.

At some later point the FS driver receives a message from Venus, namely when Venus calls sendmsg_to_kernel. At this moment the Coda FS driver looks at the contents of the message and decides if:

- the message is a reply for a suspended thread P. If so it removes the message from the processing queue and marks the message as WRITTEN. Finally, the FS driver unblocks P (still in the kernel mode context of Venus) and the sendmsg_to_kernel call returns to Venus. The process P will be scheduled at some point and continues processing its upcall with the data buffer replaced with the reply from Venus.
- The message is a `downcall`. A downcall is a request from Venus to the FS Driver. The FS driver processes the request immediately (usually a cache eviction or replacement) and when it finishes sendmsg_to_kernel returns.

Now P awakes and continues processing upcall. There are some subtleties to take account of. First P will determine if it was woken up in upcall by a signal from some other source (for example an attempt to terminate P) or as is normally the case by Venus in

its sendmsg_to_kernel call. In the normal case, the upcall routine will deallocate the message structure and return. The FS routine can proceed with its processing.

**Sleeping and IPC arrangements**

In case P is woken up by a signal and not by Venus, it will first look at the flags field. If the message is not yet READ, the process P can handle its signal without notifying Venus. If Venus has READ, and the request should not be processed, P can send Venus a signal message to indicate that it should disregard the previous message. Such signals are put in the queue at the head, and read first by Venus. If the message is already marked as WRITTEN it is too late to stop the processing. The VFS routine will now continue. (-- If a VFS request involves more than one upcall, this can lead to complicated state, an extra field ”handle_signals” could be added in the message structure to indicate points of no return have been passed.--)

### 3.1. Implementation details

The Unix implementation of this mechanism has been through the implementation of a character device associated with Coda. Venus retrieves messages by doing a read on the device, replies are sent with a write and notification is through the select system call on the file descriptor for the device. The process P is kept waiting on an interruptible wait queue object.

In Windows NT and the DPMI Windows 95 implementation a DeviceIoControl call is used. The DeviceIoControl call is designed to copy buffers from user memory to kernel memory with OPCODES. The sendmsg_to_kernel is issued as a synchronous call, while the getmsg_from_kernel call is asynchronous. Windows EventObjects are used for notification of message arrival. The process P is kept waiting on a KernelEvent object in NT and a semaphore in Windows 95.

#### 3.11.4 4. The interface at the call level

This section describes the upcalls a Coda FS driver can make to Venus. Each of these upcalls make use of two structures: inputArgs and outputArgs. In pseudo BNF form the structures take the following form:

```
struct inputArgs {
    u_long opcode;
    u_long unique;     /* Keep multiple outstanding msgs distinct */
    u_short pid;                /* Common to all */
    u_short pgid;               /* Common to all */
    struct CodaCred cred;       /* Common to all */

    <union "in" of call dependent parts of inputArgs>
};

struct outputArgs {
    u_long opcode;
    u_long unique;       /* Keep multiple outstanding msgs distinct */
    u_long result;
```

```
    <union "out" of call dependent parts of inputArgs>
};
```

Before going on let us elucidate the role of the various fields. The inputArgs start with the opcode which defines the type of service requested from Venus. There are approximately 30 upcalls at present which we will discuss. The unique field labels the inputArg with a unique number which will identify the message uniquely. A process and process group id are passed. Finally the credentials of the caller are included.

Before delving into the specific calls we need to discuss a variety of data structures shared by the kernel and Venus.

### 4.1. Data structures shared by the kernel and Venus

The CodaCred structure defines a variety of user and group ids as they are set for the calling process. The vuid_t and vgid_t are 32 bit unsigned integers. It also defines group membership in an array. On Unix the CodaCred has proven sufficient to implement good security semantics for Coda but the structure may have to undergo modification for the Windows environment when these mature:

```
struct CodaCred {
    vuid_t cr_uid, cr_euid, cr_suid, cr_fsuid; /* Real, effective, set,
↪ fs uid */
    vgid_t cr_gid, cr_egid, cr_sgid, cr_fsgid; /* same for groups */
    vgid_t cr_groups[NGROUPS];        /* Group membership for caller */
};
```

---

**Note:** It is questionable if we need CodaCreds in Venus. Finally Venus doesn't know about groups, although it does create files with the default uid/gid. Perhaps the list of group membership is superfluous.

---

The next item is the fundamental identifier used to identify Coda files, the ViceFid. A fid of a file uniquely defines a file or directory in the Coda filesystem within a cell[1]:

```
typedef struct ViceFid {
    VolumeId Volume;
    VnodeId Vnode;
    Unique_t Unique;
} ViceFid;
```

Each of the constituent fields: VolumeId, VnodeId and Unique_t are unsigned 32 bit integers. We envisage that a further field will need to be prefixed to identify the Coda cell; this will probably take the form of a Ipv6 size IP address naming the Coda cell through DNS.

The next important structure shared between Venus and the kernel is the attributes of the file. The following structure is used to exchange information. It has room for future extensions such as support for device files (currently not present in Coda):

[1] A cell is agroup of Coda servers acting under the aegis of a single system control machine or SCM. See the Coda Administration manual for a detailed description of the role of the SCM.

```
struct coda_timespec {
        int64_t         tv_sec;         /* seconds */
        long            tv_nsec;        /* nanoseconds */
};

struct coda_vattr {
        enum coda_vtype va_type;        /* vnode type (for create) */
        u_short         va_mode;        /* files access mode and type␣
↪*/
        short           va_nlink;       /* number of references to␣
↪file */
        vuid_t          va_uid;         /* owner user id */
        vgid_t          va_gid;         /* owner group id */
        long            va_fsid;        /* file system id (dev for␣
↪now) */
        long            va_fileid;      /* file id */
        u_quad_t        va_size;        /* file size in bytes */
        long            va_blocksize;   /* blocksize preferred for i/o␣
↪*/
        struct coda_timespec va_atime;  /* time of last access */
        struct coda_timespec va_mtime;  /* time of last modification */
        struct coda_timespec va_ctime;  /* time file changed */
        u_long          va_gen;         /* generation number of file */
        u_long          va_flags;       /* flags defined for file */
        dev_t           va_rdev;        /* device special file␣
↪represents */
        u_quad_t        va_bytes;       /* bytes of disk space held by␣
↪file */
        u_quad_t        va_filerev;     /* file modification number */
        u_int           va_vaflags;     /* operations flags, see below␣
↪*/
        long            va_spare;       /* remain quad aligned */
};
```

### 4.2. The pioctl interface

Coda specific requests can be made by application through the pioctl interface. The pioctl is implemented as an ordinary ioctl on a fictitious file /coda/.CONTROL. The pioctl call opens this file, gets a file handle and makes the ioctl call. Finally it closes the file.

The kernel involvement in this is limited to providing the facility to open and close and pass the ioctl message and to verify that a path in the pioctl data buffers is a file in a Coda filesystem.

The kernel is handed a data packet of the form:

```
struct {
    const char *path;
    struct ViceIoctl vidata;
```

```
    int follow;
} data;
```

where:

```
struct ViceIoctl {
        caddr_t in, out;        /* Data to be transferred in, or out */
        short in_size;          /* Size of input buffer <= 2K */
        short out_size;         /* Maximum size of output buffer, <=␣
↪2K */
};
```

The path must be a Coda file, otherwise the ioctl upcall will not be made.

---

**Note:** The data structures and code are a mess. We need to clean this up.

---

**We now proceed to document the individual calls**:

## 4.3. root

**Arguments**

in

empty

out:

```
struct cfs_root_out {
    ViceFid VFid;
} cfs_root;
```

**Description**

This call is made to Venus during the initialization of the Coda filesystem. If the result is zero, the cfs_root structure contains the ViceFid of the root of the Coda filesystem. If a non-zero result is generated, its value is a platform dependent error code indicating the difficulty Venus encountered in locating the root of the Coda filesystem.

## 4.4. lookup

**Summary**

Find the ViceFid and type of an object in a directory if it exists.

**Arguments**

in:

```
struct  cfs_lookup_in {
    ViceFid     VFid;
    char        *name;          /* Place holder for data. */
} cfs_lookup;
```

out:

```
struct cfs_lookup_out {
    ViceFid VFid;
    int vtype;
} cfs_lookup;
```

**Description**

This call is made to determine the ViceFid and filetype of a directory entry. The directory entry requested carries name 'name' and Venus will search the directory identified by cfs_lookup_in.VFid. The result may indicate that the name does not exist, or that difficulty was encountered in finding it (e.g. due to disconnection). If the result is zero, the field cfs_lookup_out.VFid contains the targets ViceFid and cfs_lookup_out.vtype the coda_vtype giving the type of object the name designates.

The name of the object is an 8 bit character string of maximum length CFS_MAXNAMLEN, currently set to 256 (including a 0 terminator.)

It is extremely important to realize that Venus bitwise ors the field cfs_lookup.vtype with CFS_NOCACHE to indicate that the object should not be put in the kernel name cache.

---

**Note:** The type of the vtype is currently wrong. It should be coda_vtype. Linux does not take note of CFS_NOCACHE. It should.

---

### 4.5. getattr

Summary Get the attributes of a file.

**Arguments**

in:

```
struct cfs_getattr_in {
    ViceFid VFid;
    struct coda_vattr attr; /* XXXXX */
} cfs_getattr;
```

out:

```
struct cfs_getattr_out {
    struct coda_vattr attr;
} cfs_getattr;
```

**Description**

This call returns the attributes of the file identified by fid.

**Errors**

Errors can occur if the object with fid does not exist, is unaccessible or if the caller does not have permission to fetch attributes.

**Note:** Many kernel FS drivers (Linux, NT and Windows 95) need to acquire the attributes as well as the Fid for the instantiation of an internal ”inode” or ”FileHandle”. A significant improvement in performance on such systems could be made by combining the lookup and getattr calls both at the Venus/kernel interaction level and at the RPC level.

The vattr structure included in the input arguments is superfluous and should be removed.

## 4.6. setattr

**Summary**

Set the attributes of a file.

**Arguments**

in:

```
struct cfs_setattr_in {
    ViceFid VFid;
    struct coda_vattr attr;
} cfs_setattr;
```

out

empty

**Description**

The structure attr is filled with attributes to be changed in BSD style. Attributes not to be changed are set to -1, apart from vtype which is set to VNON. Other are set to the value to be assigned. The only attributes which the FS driver may request to change are the mode, owner, groupid, atime, mtime and ctime. The return value indicates success or failure.

**Errors**

A variety of errors can occur. The object may not exist, may be inaccessible, or permission may not be granted by Venus.

## 4.7. access

**Arguments**

in:

```
struct cfs_access_in {
    ViceFid     VFid;
    int flags;
} cfs_access;
```

out

empty

**Description**

Verify if access to the object identified by VFid for operations described by flags is permitted. The result indicates if access will be granted. It is important to remember that Coda uses ACLs to enforce protection and that ultimately the servers, not the clients enforce the security of the system. The result of this call will depend on whether a token is held by the user.

**Errors**

The object may not exist, or the ACL describing the protection may not be accessible.

### 4.8. create

**Summary**

Invoked to create a file

**Arguments**

in:

```
struct cfs_create_in {
    ViceFid VFid;
    struct coda_vattr attr;
    int excl;
    int mode;
    char        *name;          /* Place holder for data. */
} cfs_create;
```

out:

```
struct cfs_create_out {
    ViceFid VFid;
    struct coda_vattr attr;
} cfs_create;
```

**Description**

This upcall is invoked to request creation of a file. The file will be created in the directory identified by VFid, its name will be name, and the mode will be mode. If excl is set an error will be returned if the file already exists. If the size field in attr is set to zero the file will be truncated. The uid and gid of the file are set by converting the CodaCred to a uid using a macro CRTOUID (this macro is platform dependent). Upon success the VFid and attributes of the file are returned. The Coda FS Driver will normally instantiate a vnode, inode or file handle at kernel level for the new object.

**Errors**

A variety of errors can occur. Permissions may be insufficient. If the object exists and is not a file the error EISDIR is returned under Unix.

---

**Note:** The packing of parameters is very inefficient and appears to indicate confusion between the system call creat and the VFS operation create. The VFS operation create is only called to create new objects. This create call differs from the Unix one in that it is not invoked to return a file descriptor. The truncate and exclusive options,

together with the mode, could simply be part of the mode as it is under Unix. There should be no flags argument; this is used in open (2) to return a file descriptor for READ or WRITE mode.

---

The attributes of the directory should be returned too, since the size and mtime changed.

## 4.9. mkdir

**Summary**
Create a new directory.

**Arguments**
in:

```
struct cfs_mkdir_in {
    ViceFid     VFid;
    struct coda_vattr attr;
    char        *name;          /* Place holder for data. */
} cfs_mkdir;
```

out:

```
struct cfs_mkdir_out {
    ViceFid VFid;
    struct coda_vattr attr;
} cfs_mkdir;
```

**Description**
This call is similar to create but creates a directory. Only the mode field in the input parameters is used for creation. Upon successful creation, the attr returned contains the attributes of the new directory.

**Errors**
As for create.

---

**Note:** The input parameter should be changed to mode instead of attributes.

---

The attributes of the parent should be returned since the size and mtime changes.

## 4.10. link

**Summary**
Create a link to an existing file.

**Arguments**
in:

```
struct cfs_link_in {
    ViceFid sourceFid;          /* cnode to link *to* */
```

```
    ViceFid destFid;            /* Directory in which to place␣
↪link */
    char        *tname;         /* Place holder for data. */
} cfs_link;
```

out

empty

**Description**

This call creates a link to the sourceFid in the directory identified by destFid with name tname. The source must reside in the target's parent, i.e. the source must be have parent destFid, i.e. Coda does not support cross directory hard links. Only the return value is relevant. It indicates success or the type of failure.

**Errors**

The usual errors can occur.

### 4.11. symlink

**Summary**

create a symbolic link

**Arguments**

in:

```
struct cfs_symlink_in {
    ViceFid     VFid;           /* Directory to put symlink in */
    char        *srcname;
    struct coda_vattr attr;
    char        *tname;
} cfs_symlink;
```

out

none

**Description**

Create a symbolic link. The link is to be placed in the directory identified by VFid and named tname. It should point to the pathname srcname. The attributes of the newly created object are to be set to attr.

---

**Note:** The attributes of the target directory should be returned since its size changed.

---

### 4.12. remove

**Summary**
Remove a file

**Arguments**
in:

```
struct cfs_remove_in {
    ViceFid     VFid;
    char        *name;          /* Place holder for data. */
} cfs_remove;
```

out

none

**Description**
Remove file named cfs_remove_in.name in directory identified by VFid.

---

**Note:** The attributes of the directory should be returned since its mtime and size may change.

---

### 4.13. rmdir

**Summary**
Remove a directory

**Arguments**
in:

```
struct cfs_rmdir_in {
    ViceFid     VFid;
    char        *name;          /* Place holder for data. */
} cfs_rmdir;
```

out

none

**Description**
Remove the directory with name 'name' from the directory identified by VFid.

---

**Note:** The attributes of the parent directory should be returned since its mtime and size may change.

---

### 4.14. readlink

**Summary**
Read the value of a symbolic link.

**Arguments**
in:

```
struct cfs_readlink_in {
    ViceFid VFid;
} cfs_readlink;
```

out:

```
struct cfs_readlink_out {
    int count;
    caddr_t     data;           /* Place holder for data. */
} cfs_readlink;
```

**Description**
This routine reads the contents of symbolic link identified by VFid into the buffer data. The buffer data must be able to hold any name up to CFS_MAXNAMLEN (PATH or NAM??).

**Errors**
No unusual errors.

### 4.15. open

**Summary**
Open a file.

**Arguments**
in:

```
struct cfs_open_in {
    ViceFid     VFid;
    int flags;
} cfs_open;
```

out:

```
struct cfs_open_out {
    dev_t       dev;
    ino_t       inode;
} cfs_open;
```

**Description**
This request asks Venus to place the file identified by VFid in its cache and to note that the calling process wishes to open it with flags as in open(2). The return value to the kernel differs for Unix and Windows systems. For Unix systems the Coda FS Driver is informed of the device and inode number of the container file

in the fields dev and inode. For Windows the path of the container file is returned to the kernel.

---

**Note:** Currently the cfs_open_out structure is not properly adapted to deal with the Windows case. It might be best to implement two upcalls, one to open aiming at a container file name, the other at a container file inode.

---

### 4.16. close

**Summary**

Close a file, update it on the servers.

**Arguments**

in:

```
struct cfs_close_in {
    ViceFid     VFid;
    int flags;
} cfs_close;
```

out

none

**Description**

Close the file identified by VFid.

---

**Note:** The flags argument is bogus and not used. However, Venus' code has room to deal with an execp input field, probably this field should be used to inform Venus that the file was closed but is still memory mapped for execution. There are comments about fetching versus not fetching the data in Venus vproc_vfscalls. This seems silly. If a file is being closed, the data in the container file is to be the new data. Here again the execp flag might be in play to create confusion: currently Venus might think a file can be flushed from the cache when it is still memory mapped. This needs to be understood.

---

### 4.17. ioctl

**Summary**

Do an ioctl on a file. This includes the pioctl interface.

**Arguments**

in:

```
struct cfs_ioctl_in {
    ViceFid VFid;
    int cmd;
    int len;
    int rwflag;
```

```
    char *data;                 /* Place holder for data. */
} cfs_ioctl;
```

out:

```
struct cfs_ioctl_out {
    int len;
    caddr_t     data;           /* Place holder for data. */
} cfs_ioctl;
```

**Description**

Do an ioctl operation on a file. The command, len and data arguments are filled as usual. flags is not used by Venus.

---

**Note:** Another bogus parameter. flags is not used. What is the business about PREFETCHING in the Venus code?

---

### 4.18. rename

**Summary**

Rename a fid.

**Arguments**

in:

```
struct cfs_rename_in {
    ViceFid     sourceFid;
    char        *srcname;
    ViceFid destFid;
    char        *destname;
} cfs_rename;
```

out

none

**Description**

Rename the object with name srcname in directory sourceFid to destname in destFid. It is important that the names srcname and destname are 0 terminated strings. Strings in Unix kernels are not always null terminated.

### 4.19. readdir

**Summary**
Read directory entries.

**Arguments**
in:

```
struct cfs_readdir_in {
    ViceFid     VFid;
    int count;
    int offset;
} cfs_readdir;
```

out:

```
struct cfs_readdir_out {
    int size;
    caddr_t     data;           /* Place holder for data. */
} cfs_readdir;
```

**Description**
Read directory entries from VFid starting at offset and read at most count bytes. Returns the data in data and returns the size in size.

---

**Note:** This call is not used. Readdir operations exploit container files. We will re-evaluate this during the directory revamp which is about to take place.

---

### 4.20. vget

**Summary**
instructs Venus to do an FSDB->Get.

**Arguments**
in:

```
struct cfs_vget_in {
    ViceFid VFid;
} cfs_vget;
```

out:

```
struct cfs_vget_out {
    ViceFid VFid;
    int vtype;
} cfs_vget;
```

**Description**
This upcall asks Venus to do a get operation on an fsobj labelled by VFid.

**Note:** This operation is not used. However, it is extremely useful since it can be used to deal with read/write memory mapped files. These can be ”pinned” in the Venus cache using vget and released with inactive.

### 4.21. fsync

**Summary**

Tell Venus to update the RVM attributes of a file.

**Arguments**

in:

```
struct cfs_fsync_in {
    ViceFid VFid;
} cfs_fsync;
```

out

none

**Description**

Ask Venus to update RVM attributes of object VFid. This should be called as part of kernel level fsync type calls. The result indicates if the syncing was successful.

**Note:** Linux does not implement this call. It should.

### 4.22. inactive

**Summary**

Tell Venus a vnode is no longer in use.

**Arguments**

in:

```
struct cfs_inactive_in {
    ViceFid VFid;
} cfs_inactive;
```

out

none

**Description**

This operation returns EOPNOTSUPP.

**Note:** This should perhaps be removed.

### 4.23. rdwr

**Summary**

Read or write from a file

**Arguments**

in:

```
struct cfs_rdwr_in {
    ViceFid     VFid;
    int rwflag;
    int count;
    int offset;
    int ioflag;
    caddr_t     data;           /* Place holder for data. */
} cfs_rdwr;
```

out:

```
struct cfs_rdwr_out {
    int rwflag;
    int count;
    caddr_t     data;   /* Place holder for data. */
} cfs_rdwr;
```

**Description**

This upcall asks Venus to read or write from a file.

---

**Note:** It should be removed since it is against the Coda philosophy that read/write operations never reach Venus. I have been told the operation does not work. It is not currently used.

---

### 4.24. odymount

**Summary**

Allows mounting multiple Coda ”filesystems” on one Unix mount point.

**Arguments**

in:

```
struct ody_mount_in {
    char        *name;          /* Place holder for data. */
} ody_mount;
```

out:

```
struct ody_mount_out {
    ViceFid VFid;
} ody_mount;
```

**Description**
Asks Venus to return the rootfid of a Coda system named name. The fid is returned in VFid.

---

**Note:** This call was used by David for dynamic sets. It should be removed since it causes a jungle of pointers in the VFS mounting area. It is not used by Coda proper. Call is not implemented by Venus.

---

### 4.25. ody_lookup

**Summary**
Looks up something.

**Arguments**
in

irrelevant

out

irrelevant

---

**Note:** Gut it. Call is not implemented by Venus.

---

### 4.26. ody_expand

**Summary**
expands something in a dynamic set.

**Arguments**
in

irrelevant

out

irrelevant

---

**Note:** Gut it. Call is not implemented by Venus.

---

### 4.27. prefetch

**Summary**
Prefetch a dynamic set.

Arguments

in

Not documented.

out

Not documented.

**Description**
Venus worker.cc has support for this call, although it is noted that it doesn't work. Not surprising, since the kernel does not have support for it. (ODY_PREFETCH is not a defined operation).

---

**Note:** Gut it. It isn't working and isn't used by Coda.

---

### 4.28. signal

**Summary**
Send Venus a signal about an upcall.

**Arguments**
in

none

out

not applicable.

**Description**
This is an out-of-band upcall to Venus to inform Venus that the calling process received a signal after Venus read the message from the input queue. Venus is supposed to clean up the operation.

**Errors**
No reply is given.

---

**Note:** We need to better understand what Venus needs to clean up and if it is doing this correctly. Also we need to handle multiple upcall per system call situations correctly. It would be important to know what state changes in Venus take place after an upcall for which the kernel is responsible for notifying Venus to clean up (e.g. open definitely is such a state change, but many others are maybe not).

---

### 3.11.5 5. The minicache and downcalls

The Coda FS Driver can cache results of lookup and access upcalls, to limit the frequency of upcalls. Upcalls carry a price since a process context switch needs to take place. The counterpart of caching the information is that Venus will notify the FS Driver that cached entries must be flushed or renamed.

The kernel code generally has to maintain a structure which links the internal file handles (called vnodes in BSD, inodes in Linux and FileHandles in Windows) with the ViceFid's which Venus maintains. The reason is that frequent translations back and forth are needed in order to make upcalls and use the results of upcalls. Such linking objects are called cnodes.

The current minicache implementations have cache entries which record the following:

1. the name of the file
2. the cnode of the directory containing the object
3. a list of CodaCred's for which the lookup is permitted.
4. the cnode of the object

The lookup call in the Coda FS Driver may request the cnode of the desired object from the cache, by passing its name, directory and the CodaCred's of the caller. The cache will return the cnode or indicate that it cannot be found. The Coda FS Driver must be careful to invalidate cache entries when it modifies or removes objects.

When Venus obtains information that indicates that cache entries are no longer valid, it will make a downcall to the kernel. Downcalls are intercepted by the Coda FS Driver and lead to cache invalidations of the kind described below. The Coda FS Driver does not return an error unless the downcall data could not be read into kernel memory.

#### 5.1. INVALIDATE

No information is available on this call.

#### 5.2. FLUSH

**Arguments**
None

**Summary**
Flush the name cache entirely.

**Description**
Venus issues this call upon startup and when it dies. This is to prevent stale cache information being held. Some operating systems allow the kernel name cache to be switched off dynamically. When this is done, this downcall is made.

## 5.3. PURGEUSER

### Arguments

```
struct cfs_purgeuser_out {/* CFS_PURGEUSER is a venus->kernel call␣
↪*/
    struct CodaCred cred;
} cfs_purgeuser;
```

### Description

Remove all entries in the cache carrying the Cred. This call is issued when tokens for a user expire or are flushed.

## 5.4. ZAPFILE

### Arguments

```
struct cfs_zapfile_out {  /* CFS_ZAPFILE is a venus->kernel call */
    ViceFid CodaFid;
} cfs_zapfile;
```

### Description

Remove all entries which have the (dir vnode, name) pair. This is issued as a result of an invalidation of cached attributes of a vnode.

---

**Note:** Call is not named correctly in NetBSD and Mach. The minicache zapfile routine takes different arguments. Linux does not implement the invalidation of attributes correctly.

---

## 5.5. ZAPDIR

### Arguments

```
struct cfs_zapdir_out {   /* CFS_ZAPDIR is a venus->kernel call */
    ViceFid CodaFid;
} cfs_zapdir;
```

### Description

Remove all entries in the cache lying in a directory CodaFid, and all children of this directory. This call is issued when Venus receives a callback on the directory.

## 5.6. ZAPVNODE

### Arguments

```
struct cfs_zapvnode_out { /* CFS_ZAPVNODE is a venus->kernel call␣
↪*/
    struct CodaCred cred;
    ViceFid VFid;
} cfs_zapvnode;
```

### Description

Remove all entries in the cache carrying the cred and VFid as in the arguments. This downcall is probably never issued.

## 5.7. PURGEFID

### Arguments

```
struct cfs_purgefid_out { /* CFS_PURGEFID is a venus->kernel call␣
↪*/
    ViceFid CodaFid;
} cfs_purgefid;
```

### Description

Flush the attribute for the file. If it is a dir (odd vnode), purge its children from the namecache and remove the file from the namecache.

## 5.8. REPLACE

### Summary

Replace the Fid's for a collection of names.

### Arguments

```
struct cfs_replace_out { /* cfs_replace is a venus->kernel call */
    ViceFid NewFid;
    ViceFid OldFid;
} cfs_replace;
```

### Description

This routine replaces a ViceFid in the name cache with another. It is added to allow Venus during reintegration to replace locally allocated temp fids while disconnected with global fids even when the reference counts on those fids are not zero.

### 3.11.6 6. Initialization and cleanup

This section gives brief hints as to desirable features for the Coda FS Driver at startup and upon shutdown or Venus failures. Before entering the discussion it is useful to repeat that the Coda FS Driver maintains the following data:

1. message queues
2. cnodes
3. name cache entries

   The name cache entries are entirely private to the driver, so they can easily be manipulated. The message queues will generally have clear points of initialization and destruction. The cnodes are much more delicate. User processes hold reference counts in Coda filesystems and it can be difficult to clean up the cnodes.

It can expect requests through:

1. the message subsystem
2. the VFS layer
3. pioctl interface

   Currently the pioctl passes through the VFS for Coda so we can treat these similarly.

#### 6.1. Requirements

The following requirements should be accommodated:

1. The message queues should have open and close routines. On Unix the opening of the character devices are such routines.

   - Before opening, no messages can be placed.
   - Opening will remove any old messages still pending.
   - Close will notify any sleeping processes that their upcall cannot be completed.
   - Close will free all memory allocated by the message queues.

2. At open the namecache shall be initialized to empty state.
3. Before the message queues are open, all VFS operations will fail. Fortunately this can be achieved by making sure than mounting the Coda filesystem cannot succeed before opening.
4. After closing of the queues, no VFS operations can succeed. Here one needs to be careful, since a few operations (lookup, read/write, readdir) can proceed without upcalls. These must be explicitly blocked.
5. Upon closing the namecache shall be flushed and disabled.
6. All memory held by cnodes can be freed without relying on upcalls.
7. Unmounting the file system can be done without relying on upcalls.

8. Mounting the Coda filesystem should fail gracefully if Venus cannot get the rootfid or the attributes of the rootfid. The latter is best implemented by Venus fetching these objects before attempting to mount.

---

**Note:** NetBSD in particular but also Linux have not implemented the above requirements fully. For smooth operation this needs to be corrected.

---

## 3.12 Configfs - Userspace-driven Kernel Object Configuration

Joel Becker <joel.becker@oracle.com>

Updated: 31 March 2005



### 3.12.1 What is configfs?

configfs is a ram-based filesystem that provides the converse of sysfs's functionality. Where sysfs is a filesystem-based view of kernel objects, configfs is a filesystem-based manager of kernel objects, or config_items.

With sysfs, an object is created in kernel (for example, when a device is discovered) and it is registered with sysfs. Its attributes then appear in sysfs, allowing userspace to read the attributes via readdir(3)/read(2). It may allow some attributes to be modified via write(2). The important point is that the object is created and destroyed in kernel, the kernel controls the lifecycle of the sysfs representation, and sysfs is merely a window on all this.

A configfs config_item is created via an explicit userspace operation: mkdir(2). It is destroyed via rmdir(2). The attributes appear at mkdir(2) time, and can be read or modified via read(2) and write(2). As with sysfs, readdir(3) queries the list of items and/or attributes. symlink(2) can be used to group items together. Unlike sysfs, the lifetime of the representation is completely driven by userspace. The kernel modules backing the items must respond to this.

Both sysfs and configfs can and should exist together on the same system. One is not a replacement for the other.

### 3.12.2 Using configfs

configfs can be compiled as a module or into the kernel. You can access it by doing:

```
mount -t configfs none /config
```

The configfs tree will be empty unless client modules are also loaded. These are modules that register their item types with configfs as subsystems. Once a client subsystem is loaded, it will appear as a subdirectory (or more than one) under /config. Like sysfs, the configfs tree is always there, whether mounted on /config or not.

An item is created via mkdir(2). The item's attributes will also appear at this time. readdir(3) can determine what the attributes are, read(2) can query their default values, and write(2) can store new values. Don't mix more than one attribute in one attribute file.

There are two types of configfs attributes:

- Normal attributes, which similar to sysfs attributes, are small ASCII text files, with a maximum size of one page (PAGE_SIZE, 4096 on i386). Preferably only one value per file should be used, and the same caveats from sysfs apply. Configfs expects write(2) to store the entire buffer at once. When writing to normal configfs attributes, userspace processes should first read the entire file, modify the portions they wish to change, and then write the entire buffer back.
- Binary attributes, which are somewhat similar to sysfs binary attributes, but with a few slight changes to semantics. The PAGE_SIZE limitation does not apply, but the whole binary item must fit in single kernel vmalloc'ed buffer. The write(2) calls from user space are buffered, and the attributes' write_bin_attribute method will be invoked on the final close, therefore it is imperative for user-space to check the return code of close(2) in order to verify that the operation finished successfully. To avoid a malicious user OOMing the kernel, there's a per-binary attribute maximum buffer value.

When an item needs to be destroyed, remove it with rmdir(2). An item cannot be destroyed if any other item has a link to it (via symlink(2)). Links can be removed via unlink(2).

### 3.12.3 Configuring FakeNBD: an Example

Imagine there's a Network Block Device (NBD) driver that allows you to access remote block devices. Call it FakeNBD. FakeNBD uses configfs for its configuration. Obviously, there will be a nice program that sysadmins use to configure FakeNBD, but somehow that program has to tell the driver about it. Here's where configfs comes in.

When the FakeNBD driver is loaded, it registers itself with configfs. readdir(3) sees this just fine:

```
# ls /config
fakenbd
```

A fakenbd connection can be created with mkdir(2). The name is arbitrary, but likely the tool will make some use of the name. Perhaps it is a uuid or a disk name:

```
# mkdir /config/fakenbd/disk1
# ls /config/fakenbd/disk1
target device rw
```

The target attribute contains the IP address of the server FakeNBD will connect to. The device attribute is the device on the server. Predictably, the rw attribute determines whether the connection is read-only or read-write:

```
# echo 10.0.0.1 > /config/fakenbd/disk1/target
# echo /dev/sda1 > /config/fakenbd/disk1/device
# echo 1 > /config/fakenbd/disk1/rw
```

That's it. That's all there is. Now the device is configured, via the shell no less.

### 3.12.4 Coding With configfs

Every object in configfs is a config_item. A config_item reflects an object in the subsystem. It has attributes that match values on that object. configfs handles the filesystem representation of that object and its attributes, allowing the subsystem to ignore all but the basic show/store interaction.

Items are created and destroyed inside a config_group. A group is a collection of items that share the same attributes and operations. Items are created by mkdir(2) and removed by rmdir(2), but configfs handles that. The group has a set of operations to perform these tasks

A subsystem is the top level of a client module. During initialization, the client module registers the subsystem with configfs, the subsystem appears as a directory at the top of the configfs filesystem. A subsystem is also a config_group, and can do everything a config_group can.

### 3.12.5 struct config_item

```
struct config_item {
        char                    *ci_name;
        char                    ci_namebuf[UOBJ_NAME_LEN];
        struct kref             ci_kref;
        struct list_head        ci_entry;
        struct config_item      *ci_parent;
        struct config_group     *ci_group;
        struct config_item_type *ci_type;
        struct dentry           *ci_dentry;
};

void config_item_init(struct config_item *);
void config_item_init_type_name(struct config_item *,
                                const char *name,
                                struct config_item_type *type);
struct config_item *config_item_get(struct config_item *);
void config_item_put(struct config_item *);
```

Generally, struct config_item is embedded in a container structure, a structure that actually represents what the subsystem is doing. The config_item portion of that structure is how the object interacts with configfs.

Whether statically defined in a source file or created by a parent config_group, a config_item must have one of the _init() functions called on it. This initializes the reference count and sets up the appropriate fields.

All users of a config_item should have a reference on it via config_item_get(), and drop the reference when they are done via config_item_put().

By itself, a config_item cannot do much more than appear in configfs. Usually a subsystem wants the item to display and/or store attributes, among other things. For that, it needs a type.

### 3.12.6 struct config_item_type

```
struct configfs_item_operations {
        void (*release)(struct config_item *);
        int (*allow_link)(struct config_item *src,
                          struct config_item *target);
        void (*drop_link)(struct config_item *src,
                          struct config_item *target);
};

struct config_item_type {
        struct module                           *ct_owner;
        struct configfs_item_operations         *ct_item_ops;
        struct configfs_group_operations        *ct_group_ops;
        struct configfs_attribute               **ct_attrs;
        struct configfs_bin_attribute           **ct_bin_attrs;
};
```

The most basic function of a config_item_type is to define what operations can be performed on a config_item. All items that have been allocated dynamically will need to provide the ct_item_ops->release() method. This method is called when the config_item's reference count reaches zero.

### 3.12.7 struct configfs_attribute

```
struct configfs_attribute {
        char                    *ca_name;
        struct module           *ca_owner;
        umode_t                  ca_mode;
        ssize_t (*show)(struct config_item *, char *);
        ssize_t (*store)(struct config_item *, const char *, size_t);
};
```

When a config_item wants an attribute to appear as a file in the item's configfs directory, it must define a configfs_attribute describing it. It then adds the attribute to the NULL-terminated array config_item_type->ct_attrs. When the item appears in configfs, the attribute file will appear with the configfs_attribute->ca_name filename. configfs_attribute->ca_mode specifies the file permissions.

If an attribute is readable and provides a ->show method, that method will be called whenever userspace asks for a read(2) on the attribute. If an attribute is writable and provides a ->store method, that method will be called whenever userspace asks for a write(2) on the attribute.

### 3.12.8 struct configfs_bin_attribute

```
struct configfs_bin_attribute {
        struct configfs_attribute       cb_attr;
        void                            *cb_private;
        size_t                          cb_max_size;
};
```

The binary attribute is used when the one needs to use binary blob to appear as the contents of a file in the item's configfs directory. To do so add the binary attribute to the NULL-terminated array config_item_type->ct_bin_attrs, and the item appears in configfs, the attribute file will appear with the configfs_bin_attribute->cb_attr.ca_name filename. configfs_bin_attribute->cb_attr.ca_mode specifies the file permissions. The cb_private member is provided for use by the driver, while the cb_max_size member specifies the maximum amount of vmalloc buffer to be used.

If binary attribute is readable and the config_item provides a ct_item_ops->read_bin_attribute() method, that method will be called whenever userspace asks for a read(2) on the attribute. The converse will happen for write(2). The reads/writes are buffered so only a single read/write will occur; the attributes' need not concern itself with it.

### 3.12.9 struct config_group

A config_item cannot live in a vacuum. The only way one can be created is via mkdir(2) on a config_group. This will trigger creation of a child item:

```
struct config_group {
        struct config_item              cg_item;
        struct list_head                cg_children;
        struct configfs_subsystem       *cg_subsys;
        struct list_head                default_groups;
        struct list_head                group_entry;
};

void config_group_init(struct config_group *group);
void config_group_init_type_name(struct config_group *group,
                                 const char *name,
                                 struct config_item_type *type);
```

The config_group structure contains a config_item. Properly configuring that item means that a group can behave as an item in its own right. However, it can do more: it can create child items or groups. This is accomplished via the group operations specified on the group's config_item_type:

```
struct configfs_group_operations {
        struct config_item *(*make_item)(struct config_group *group,
                                         const char *name);
        struct config_group *(*make_group)(struct config_group *group,
                                           const char *name);
        void (*disconnect_notify)(struct config_group *group,
                                  struct config_item *item);
```

```
        void (*drop_item)(struct config_group *group,
                          struct config_item *item);
};
```

A group creates child items by providing the ct_group_ops->make_item() method. If provided, this method is called from mkdir(2) in the group's directory. The subsystem allocates a new config_item (or more likely, its container structure), initializes it, and returns it to configfs. Configfs will then populate the filesystem tree to reflect the new item.

If the subsystem wants the child to be a group itself, the subsystem provides ct_group_ops->make_group(). Everything else behaves the same, using the group _init() functions on the group.

Finally, when userspace calls rmdir(2) on the item or group, ct_group_ops->drop_item() is called. As a config_group is also a config_item, it is not necessary for a separate drop_group() method. The subsystem must config_item_put() the reference that was initialized upon item allocation. If a subsystem has no work to do, it may omit the ct_group_ops->drop_item() method, and configfs will call config_item_put() on the item on behalf of the subsystem.

**Important:**

drop_item() is void, and as such cannot fail. When rmdir(2) is called, configfs WILL remove the item from the filesystem tree (assuming that it has no children to keep it busy). The subsystem is responsible for responding to this. If the subsystem has references to the item in other threads, the memory is safe. It may take some time for the item to actually disappear from the subsystem's usage. But it is gone from configfs.

When drop_item() is called, the item's linkage has already been torn down. It no longer has a reference on its parent and has no place in the item hierarchy. If a client needs to do some cleanup before this teardown happens, the subsystem can implement the ct_group_ops->disconnect_notify() method. The method is called after configfs has removed the item from the filesystem view but before the item is removed from its parent group. Like drop_item(), disconnect_notify() is void and cannot fail. Client subsystems should not drop any references here, as they still must do it in drop_item().

A config_group cannot be removed while it still has child items. This is implemented in the configfs rmdir(2) code. ->drop_item() will not be called, as the item has not been dropped. rmdir(2) will fail, as the directory is not empty.

### 3.12.10 struct configfs_subsystem

A subsystem must register itself, usually at module_init time. This tells configfs to make the subsystem appear in the file tree:

```
struct configfs_subsystem {
        struct config_group     su_group;
        struct mutex            su_mutex;
};

int configfs_register_subsystem(struct configfs_subsystem *subsys);
void configfs_unregister_subsystem(struct configfs_subsystem *subsys);
```

A subsystem consists of a toplevel config_group and a mutex. The group is where child config_items are created. For a subsystem, this group is usually defined statically. Before call-

ing configfs_register_subsystem(), the subsystem must have initialized the group via the usual group _init() functions, and it must also have initialized the mutex.

When the register call returns, the subsystem is live, and it will be visible via configfs. At that point, mkdir(2) can be called and the subsystem must be ready for it.

### 3.12.11 An Example

The best example of these basic concepts is the simple_children subsystem/group and the simple_child item in samples/configfs/configfs_sample.c. It shows a trivial object displaying and storing an attribute, and a simple group creating and destroying these children.

### 3.12.12 Hierarchy Navigation and the Subsystem Mutex

There is an extra bonus that configfs provides. The config_groups and config_items are arranged in a hierarchy due to the fact that they appear in a filesystem. A subsystem is NEVER to touch the filesystem parts, but the subsystem might be interested in this hierarchy. For this reason, the hierarchy is mirrored via the config_group->cg_children and config_item->ci_parent structure members.

A subsystem can navigate the cg_children list and the ci_parent pointer to see the tree created by the subsystem. This can race with configfs' management of the hierarchy, so configfs uses the subsystem mutex to protect modifications. Whenever a subsystem wants to navigate the hierarchy, it must do so under the protection of the subsystem mutex.

A subsystem will be prevented from acquiring the mutex while a newly allocated item has not been linked into this hierarchy. Similarly, it will not be able to acquire the mutex while a dropping item has not yet been unlinked. This means that an item's ci_parent pointer will never be NULL while the item is in configfs, and that an item will only be in its parent's cg_children list for the same duration. This allows a subsystem to trust ci_parent and cg_children while they hold the mutex.

### 3.12.13 Item Aggregation Via symlink(2)

configfs provides a simple group via the group->item parent/child relationship. Often, however, a larger environment requires aggregation outside of the parent/child connection. This is implemented via symlink(2).

A config_item may provide the ct_item_ops->allow_link() and ct_item_ops->drop_link() methods. If the ->allow_link() method exists, symlink(2) may be called with the config_item as the source of the link. These links are only allowed between configfs config_items. Any symlink(2) attempt outside the configfs filesystem will be denied.

When symlink(2) is called, the source config_item's ->allow_link() method is called with itself and a target item. If the source item allows linking to target item, it returns 0. A source item may wish to reject a link if it only wants links to a certain type of object (say, in its own subsystem).

When unlink(2) is called on the symbolic link, the source item is notified via the ->drop_link() method. Like the ->drop_item() method, this is a void function and cannot return failure. The subsystem is responsible for responding to the change.

A config_item cannot be removed while it links to any other item, nor can it be removed while an item links to it. Dangling symlinks are not allowed in configfs.

### 3.12.14 Automatically Created Subgroups

A new config_group may want to have two types of child config_items. While this could be codified by magic names in ->make_item(), it is much more explicit to have a method whereby userspace sees this divergence.

Rather than have a group where some items behave differently than others, configfs provides a method whereby one or many subgroups are automatically created inside the parent at its creation. Thus, mkdir("parent") results in "parent", "parent/subgroup1", up through "parent/subgroupN". Items of type 1 can now be created in "parent/subgroup1", and items of type N can be created in "parent/subgroupN".

These automatic subgroups, or default groups, do not preclude other children of the parent group. If ct_group_ops->make_group() exists, other child groups can be created on the parent group directly.

A configfs subsystem specifies default groups by adding them using the configfs_add_default_group() function to the parent config_group structure. Each added group is populated in the configfs tree at the same time as the parent group. Similarly, they are removed at the same time as the parent. No extra notification is provided. When a ->drop_item() method call notifies the subsystem the parent group is going away, it also means every default group child associated with that parent group.

As a consequence of this, default groups cannot be removed directly via rmdir(2). They also are not considered when rmdir(2) on the parent group is checking for children.

### 3.12.15 Dependent Subsystems

Sometimes other drivers depend on particular configfs items. For example, ocfs2 mounts depend on a heartbeat region item. If that region item is removed with rmdir(2), the ocfs2 mount must BUG or go readonly. Not happy.

configfs provides two additional API calls: configfs_depend_item() and configfs_undepend_item(). A client driver can call configfs_depend_item() on an existing item to tell configfs that it is depended on. configfs will then return -EBUSY from rmdir(2) for that item. When the item is no longer depended on, the client driver calls configfs_undepend_item() on it.

These API cannot be called underneath any configfs callbacks, as they will conflict. They can block and allocate. A client driver probably shouldn't calling them of its own gumption. Rather it should be providing an API that external subsystems call.

How does this work? Imagine the ocfs2 mount process. When it mounts, it asks for a heartbeat region item. This is done via a call into the heartbeat code. Inside the heartbeat code, the region item is looked up. Here, the heartbeat code calls configfs_depend_item(). If it succeeds, then heartbeat knows the region is safe to give to ocfs2. If it fails, it was being torn down anyway, and heartbeat can gracefully pass up an error.

## 3.13 Cramfs - cram a filesystem onto a small ROM

cramfs is designed to be simple and small, and to compress things well.

It uses the zlib routines to compress a file one page at a time, and allows random page access. The meta-data is not compressed, but is expressed in a very terse representation to make it use much less diskspace than traditional filesystems.

You can't write to a cramfs filesystem (making it compressible and compact also makes it _very_ hard to update on-the-fly), so you have to create the disk image with the ”mkcramfs” utility.

### 3.13.1 Usage Notes

File sizes are limited to less than 16MB.

Maximum filesystem size is a little over 256MB. (The last file on the filesystem is allowed to extend past 256MB.)

Only the low 8 bits of gid are stored. The current version of mkcramfs simply truncates to 8 bits, which is a potential security issue.

Hard links are supported, but hard linked files will still have a link count of 1 in the cramfs image.

Cramfs directories have no `.` or `..` entries. Directories (like every other file on cramfs) always have a link count of 1. (There's no need to use -noleaf in `find`, btw.)

No timestamps are stored in a cramfs, so these default to the epoch (1970 GMT). Recently-accessed files may have updated timestamps, but the update lasts only as long as the inode is cached in memory, after which the timestamp reverts to 1970, i.e. moves backwards in time.

Currently, cramfs must be written and read with architectures of the same endianness, and can be read only by kernels with PAGE_SIZE == 4096. At least the latter of these is a bug, but it hasn't been decided what the best fix is. For the moment if you have larger pages you can just change the #define in mkcramfs.c, so long as you don't mind the filesystem becoming unreadable to future kernels.

### 3.13.2 Memory Mapped cramfs image

The CRAMFS_MTD Kconfig option adds support for loading data directly from a physical linear memory range (usually non volatile memory like Flash) instead of going through the block device layer. This saves some memory since no intermediate buffering is necessary to hold the data before decompressing.

And when data blocks are kept uncompressed and properly aligned, they will automatically be mapped directly into user space whenever possible providing eXecute-In-Place (XIP) from ROM of read-only segments. Data segments mapped read-write (hence they have to be copied to RAM) may still be compressed in the cramfs image in the same file along with non compressed read-only segments. Both MMU and no-MMU systems are supported. This is particularly handy for tiny embedded systems with very tight memory constraints.

The location of the cramfs image in memory is system dependent. You must know the proper physical address where the cramfs image is located and configure an MTD device for it. Also,

that MTD device must be supported by a map driver that implements the ”point” method. Examples of such MTD drivers are cfi_cmdset_0001 (Intel/Sharp CFI flash) or physmap (Flash device in physical memory map). MTD partitions based on such devices are fine too. Then that device should be specified with the ”mtd:” prefix as the mount device argument. For example, to mount the MTD device named ”fs_partition” on the /mnt directory:

```
$ mount -t cramfs mtd:fs_partition /mnt
```

To boot a kernel with this as root filesystem, suffice to specify something like ”root=mtd:fs_partition” on the kernel command line.

### 3.13.3 Tools

A version of mkcramfs that can take advantage of the latest capabilities described above can be found here:

https://github.com/npitre/cramfs-tools

### 3.13.4 For /usr/share/magic

| | | |
|---|---|---|
| 0 | ulelong 0x28cd3d45 | Linux cramfs offset 0 |
| >4 | ulelong x | size %d |
| >8 | ulelong x | flags 0x%x |
| >12 | ulelong x | future 0x%x |
| >16 | string >0 | signature ”%.16s” |
| >32 | ulelong x | fsid.crc 0x%x |
| >36 | ulelong x | fsid.edition %d |
| >40 | ulelong x | fsid.blocks %d |
| >44 | ulelong x | fsid.files %d |
| >48 | string >0 | name ”%.16s” |
| 512 | ulelong 0x28cd3d45 | Linux cramfs offset 512 |
| >516 | ulelong x | size %d |
| >520 | ulelong x | flags 0x%x |
| >524 | ulelong x | future 0x%x |
| >528 | string >0 | signature ”%.16s” |
| >544 | ulelong x | fsid.crc 0x%x |
| >548 | ulelong x | fsid.edition %d |
| >552 | ulelong x | fsid.blocks %d |
| >556 | ulelong x | fsid.files %d |
| >560 | string >0 | name ”%.16s” |

### 3.13.5 Hacker Notes

See fs/cramfs/README for filesystem layout and implementation notes.

## 3.14 Direct Access for files

### 3.14.1 Motivation

The page cache is usually used to buffer reads and writes to files. It is also used to provide the pages which are mapped into userspace by a call to mmap.

For block devices that are memory-like, the page cache pages would be unnecessary copies of the original storage. The *DAX* code removes the extra copy by performing reads and writes directly to the storage device. For file mappings, the storage device is mapped directly into userspace.

### 3.14.2 Usage

If you have a block device which supports *DAX*, you can make a filesystem on it as usual. The *DAX* code currently only supports files with a block size equal to your kernel's *PAGE_SIZE*, so you may need to specify a block size when creating the filesystem.

Currently 5 filesystems support *DAX*: ext2, ext4, xfs, virtiofs and erofs. Enabling *DAX* on them is different.

### 3.14.3 Enabling DAX on ext2 and erofs

When mounting the filesystem, use the `-o dax` option on the command line or add 'dax' to the options in `/etc/fstab`. This works to enable *DAX* on all files within the filesystem. It is equivalent to the `-o dax=always` behavior below.

### 3.14.4 Enabling DAX on xfs and ext4

### 3.14.5 Summary

1. There exists an in-kernel file access mode flag *S_DAX* that corresponds to the statx flag *STATX_ATTR_DAX*. See the manpage for statx(2) for details about this access mode.
2. There exists a persistent flag *FS_XFLAG_DAX* that can be applied to regular files and directories. This advisory flag can be set or cleared at any time, but doing so does not immediately affect the *S_DAX* state.
3. If the persistent *FS_XFLAG_DAX* flag is set on a directory, this flag will be inherited by all regular files and subdirectories that are subsequently created in this directory. Files and subdirectories that exist at the time this flag is set or cleared on the parent directory are not modified by this modification of the parent directory.
4. There exist dax mount options which can override *FS_XFLAG_DAX* in the setting of the *S_DAX* flag. Given underlying storage which supports *DAX* the following hold:

   `-o dax=inode` means ”follow *FS_XFLAG_DAX*” and is the default.

`-o dax=never` means "never set *S_DAX*, ignore *FS_XFLAG_DAX*."

`-o dax=always` means "always set *S_DAX* ignore *FS_XFLAG_DAX*."

`-o dax` is a legacy option which is an alias for `dax=always`.

> **Warning:** The option `-o dax` may be removed in the future so `-o dax=always` is the preferred method for specifying this behavior.

---

**Note:** Modifications to and the inheritance behavior of *FS_XFLAG_DAX* remain the same even when the filesystem is mounted with a dax option. However, in-core inode state (*S_DAX*) will be overridden until the filesystem is remounted with dax=inode and the inode is evicted from kernel memory.

---

5. The *S_DAX* policy can be changed via:
   a) Setting the parent directory *FS_XFLAG_DAX* as needed before files are created
   b) Setting the appropriate dax="foo" mount option
   c) Changing the *FS_XFLAG_DAX* flag on existing regular files and directories. This has runtime constraints and limitations that are described in 6) below.
6. When changing the *S_DAX* policy via toggling the persistent *FS_XFLAG_DAX* flag, the change to existing regular files won't take effect until the files are closed by all processes.

### 3.14.6 Details

There are 2 per-file dax flags. One is a persistent inode setting (*FS_XFLAG_DAX*) and the other is a volatile flag indicating the active state of the feature (*S_DAX*).

*FS_XFLAG_DAX* is preserved within the filesystem. This persistent config setting can be set, cleared and/or queried using the *FS_IOC_FS`[`GS*]`ETXATTR` ioctl (see ioctl_xfs_fsgetxattr(2)) or an utility such as 'xfs_io'.

New files and directories automatically inherit *FS_XFLAG_DAX* from their parent directory **when created**. Therefore, setting *FS_XFLAG_DAX* at directory creation time can be used to set a default behavior for an entire sub-tree.

To clarify inheritance, here are 3 examples:

Example A:

```
mkdir -p a/b/c
xfs_io -c 'chattr +x' a
mkdir a/b/c/d
mkdir a/e

------[outcome]------

dax: a,e
no dax: b,c,d
```

Example B:

```
mkdir a
xfs_io -c 'chattr +x' a
mkdir -p a/b/c/d

------[outcome]------

dax: a,b,c,d
no dax:
```

Example C:

```
mkdir -p a/b/c
xfs_io -c 'chattr +x' c
mkdir a/b/c/d

------[outcome]------

dax: c,d
no dax: a,b
```

The current enabled state (*S_DAX*) is set when a file inode is instantiated in memory by the kernel. It is set based on the underlying media support, the value of *FS_XFLAG_DAX* and the filesystem's dax mount option.

statx can be used to query *S_DAX*.

---

**Note:** That only regular files will ever have *S_DAX* set and therefore statx will never indicate that *S_DAX* is set on directories.

---

Setting the *FS_XFLAG_DAX* flag (specifically or through inheritance) occurs even if the underlying media does not support dax and/or the filesystem is overridden with a mount option.

### 3.14.7 Enabling DAX on virtiofs

The semantic of DAX on virtiofs is basically equal to that on ext4 and xfs, except that when '-o dax=inode' is specified, virtiofs client derives the hint whether DAX shall be enabled or not from virtiofs server through FUSE protocol, rather than the persistent *FS_XFLAG_DAX* flag. That is, whether DAX shall be enabled or not is completely determined by virtiofs server, while virtiofs server itself may deploy various algorithm making this decision, e.g. depending on the persistent *FS_XFLAG_DAX* flag on the host.

It is still supported to set or clear persistent *FS_XFLAG_DAX* flag inside guest, but it is not guaranteed that DAX will be enabled or disabled for corresponding file then. Users inside guest still need to call statx(2) and check the statx flag *STATX_ATTR_DAX* to see if DAX is enabled for this file.

### 3.14.8 Implementation Tips for Block Driver Writers

To support *DAX* in your block driver, implement the 'direct_access' block device operation. It is used to translate the sector number (expressed in units of 512-byte sectors) to a page frame number (pfn) that identifies the physical page for the memory. It also returns a kernel virtual address that can be used to access the memory.

The direct_access method takes a 'size' parameter that indicates the number of bytes being requested. The function should return the number of bytes that can be contiguously accessed at that offset. It may also return a negative errno if an error occurs.

In order to support this method, the storage must be byte-accessible by the CPU at all times. If your device uses paging techniques to expose a large amount of memory through a smaller window, then you cannot implement direct_access. Equally, if your device can occasionally stall the CPU for an extended period, you should also not attempt to implement direct_access.

These block devices may be used for inspiration: - brd: RAM backed block device driver - dcssblk: s390 dcss block device driver - pmem: NVDIMM persistent memory driver

### 3.14.9 Implementation Tips for Filesystem Writers

Filesystem support consists of:

- Adding support to mark inodes as being *DAX* by setting the *S_DAX* flag in i_flags
- Implementing ->read_iter and ->write_iter operations which use *dax_iomap_rw()* when inode has *S_DAX* flag set
- Implementing an mmap file operation for *DAX* files which sets the *VM_MIXEDMAP* and *VM_HUGEPAGE* flags on the *VMA*, and setting the vm_ops to include handlers for fault, pmd_fault, page_mkwrite, pfn_mkwrite. These handlers should probably call *dax_iomap_fault()* passing the appropriate fault size and iomap operations.
- Calling `iomap_zero_range()` passing appropriate iomap operations instead of `block_truncate_page()` for *DAX* files
- Ensuring that there is sufficient locking between reads, writes, truncates and page faults

The iomap handlers for allocating blocks must make sure that allocated blocks are zeroed out and converted to written extents before being returned to avoid exposure of uninitialized data through mmap.

These filesystems may be used for inspiration:

**See also:**

ext2: see *The Second Extended Filesystem*

**See also:**

xfs: see Documentation/admin-guide/xfs.rst

**See also:**

ext4: see Documentation/filesystems/ext4/

### 3.14.10 Handling Media Errors

The libnvdimm subsystem stores a record of known media error locations for each pmem block device (in gendisk->badblocks). If we fault at such location, or one with a latent error not yet discovered, the application can expect to receive a *SIGBUS*. Libnvdimm also allows clearing of these errors by simply writing the affected sectors (through the pmem driver, and if the underlying NVDIMM supports the clear_poison DSM defined by ACPI).

Since *DAX* IO normally doesn't go through the `driver/bio` path, applications or sysadmins have an option to restore the lost data from a prior `backup/inbuilt` redundancy in the following ways:

1. Delete the affected file, and restore from a backup (sysadmin route): This will free the filesystem blocks that were being used by the file, and the next time they're allocated, they will be zeroed first, which happens through the driver, and will clear bad sectors.
2. Truncate or hole-punch the part of the file that has a bad-block (at least an entire aligned sector has to be hole-punched, but not necessarily an entire filesystem block).

These are the two basic paths that allow *DAX* filesystems to continue operating in the presence of media errors. More robust error recovery mechanisms can be built on top of this in the future, for example, involving redundancy/mirroring provided at the block layer through DM, or additionally, at the filesystem level. These would have to rely on the above two tenets, that error clearing can happen either by sending an IO through the driver, or zeroing (also through the driver).

### 3.14.11 Shortcomings

Even if the kernel or its modules are stored on a filesystem that supports *DAX* on a block device that supports *DAX*, they will still be copied into RAM.

The DAX code does not work correctly on architectures which have virtually mapped caches such as ARM, MIPS and SPARC.

Calling `get_user_pages()` on a range of user memory that has been mmapped from a *DAX* file will fail when there are no 'struct page' to describe those pages. This problem has been addressed in some device drivers by adding optional struct page support for pages under the control of the driver (see *CONFIG_NVDIMM_PFN* in `drivers/nvdimm` for an example of how to do this). In the non struct page cases *O_DIRECT* reads/writes to those memory ranges from a non-*DAX* file will fail

---

**Note:** *O_DIRECT* reads/writes _of a *DAX* file do work, it is the memory that is being accessed that is key here). Other things that will not work in the non struct page case include RDMA, `sendfile()` and `splice()`.

---

## 3.15 DebugFS



Debugfs exists as a simple way for kernel developers to make information available to user space. Unlike /proc, which is only meant for information about a process, or sysfs, which has strict one-value-per-file rules, debugfs has no rules at all. Developers can put any information they want there. The debugfs filesystem is also intended to not serve as a stable ABI to user space; in theory, there are no stability constraints placed on files exported there. The real world is not always so simple, though[1]; even debugfs interfaces are best designed with the idea that they will need to be maintained forever.

Debugfs is typically mounted with a command like:

```
mount -t debugfs none /sys/kernel/debug
```

(Or an equivalent /etc/fstab line). The debugfs root directory is accessible only to the root user by default. To change access to the tree the "uid", "gid" and "mode" mount options can be used.

Note that the debugfs API is exported GPL-only to modules.

Code using debugfs should include <linux/debugfs.h>. Then, the first order of business will be to create at least one directory to hold a set of debugfs files:

```
struct dentry *debugfs_create_dir(const char *name, struct dentry *parent);
```

This call, if successful, will make a directory called name underneath the indicated parent directory. If parent is NULL, the directory will be created in the debugfs root. On success, the return value is a struct dentry pointer which can be used to create files in the directory (and to clean it up at the end). An ERR_PTR(-ERROR) return value indicates that something went wrong. If ERR_PTR(-ENODEV) is returned, that is an indication that the kernel has been built without debugfs support and none of the functions described below will work.

The most general way to create a file within a debugfs directory is with:

```
struct dentry *debugfs_create_file(const char *name, umode_t mode,
                                   struct dentry *parent, void *data,
                                   const struct file_operations *fops);
```

Here, name is the name of the file to create, mode describes the access permissions the file should have, parent indicates the directory which should hold the file, data will be stored in the i_private field of the resulting inode structure, and fops is a set of file operations which implement the file's behavior. At a minimum, the read() and/or write() operations should be provided; others can be included as needed. Again, the return value will be a dentry pointer to the created file, ERR_PTR(-ERROR) on error, or ERR_PTR(-ENODEV) if debugfs support is missing.

Create a file with an initial size, the following function can be used instead:

```
void debugfs_create_file_size(const char *name, umode_t mode,
                              struct dentry *parent, void *data,
                              const struct file_operations *fops,
                              loff_t file_size);
```

[1] http://lwn.net/Articles/309298/

file_size is the initial file size. The other parameters are the same as the function debugfs_create_file.

In a number of cases, the creation of a set of file operations is not actually necessary; the debugfs code provides a number of helper functions for simple situations. Files containing a single integer value can be created with any of:

```
void debugfs_create_u8(const char *name, umode_t mode,
                       struct dentry *parent, u8 *value);
void debugfs_create_u16(const char *name, umode_t mode,
                        struct dentry *parent, u16 *value);
void debugfs_create_u32(const char *name, umode_t mode,
                        struct dentry *parent, u32 *value);
void debugfs_create_u64(const char *name, umode_t mode,
                        struct dentry *parent, u64 *value);
```

These files support both reading and writing the given value; if a specific file should not be written to, simply set the mode bits accordingly. The values in these files are in decimal; if hexadecimal is more appropriate, the following functions can be used instead:

```
void debugfs_create_x8(const char *name, umode_t mode,
                       struct dentry *parent, u8 *value);
void debugfs_create_x16(const char *name, umode_t mode,
                        struct dentry *parent, u16 *value);
void debugfs_create_x32(const char *name, umode_t mode,
                        struct dentry *parent, u32 *value);
void debugfs_create_x64(const char *name, umode_t mode,
                        struct dentry *parent, u64 *value);
```

These functions are useful as long as the developer knows the size of the value to be exported. Some types can have different widths on different architectures, though, complicating the situation somewhat. There are functions meant to help out in such special cases:

```
void debugfs_create_size_t(const char *name, umode_t mode,
                           struct dentry *parent, size_t *value);
```

As might be expected, this function will create a debugfs file to represent a variable of type size_t.

Similarly, there are helpers for variables of type unsigned long, in decimal and hexadecimal:

```
struct dentry *debugfs_create_ulong(const char *name, umode_t mode,
                                    struct dentry *parent,
                                    unsigned long *value);
void debugfs_create_xul(const char *name, umode_t mode,
                        struct dentry *parent, unsigned long *value);
```

Boolean values can be placed in debugfs with:

```
void debugfs_create_bool(const char *name, umode_t mode,
                         struct dentry *parent, bool *value);
```

A read on the resulting file will yield either Y (for non-zero values) or N, followed by a newline. If written to, it will accept either upper- or lower-case values, or 1 or 0. Any other input will be

silently ignored.

Also, atomic_t values can be placed in debugfs with:

```
void debugfs_create_atomic_t(const char *name, umode_t mode,
                             struct dentry *parent, atomic_t *value)
```

A read of this file will get atomic_t values, and a write of this file will set atomic_t values.

Another option is exporting a block of arbitrary binary data, with this structure and function:

```
struct debugfs_blob_wrapper {
    void *data;
    unsigned long size;
};

struct dentry *debugfs_create_blob(const char *name, umode_t mode,
                                   struct dentry *parent,
                                   struct debugfs_blob_wrapper *blob);
```

A read of this file will return the data pointed to by the debugfs_blob_wrapper structure. Some drivers use ”blobs” as a simple way to return several lines of (static) formatted text output. This function can be used to export binary information, but there does not appear to be any code which does so in the mainline. Note that all files created with *debugfs_create_blob()* are read-only.

If you want to dump a block of registers (something that happens quite often during development, even if little such code reaches mainline), debugfs offers two functions: one to make a registers-only file, and another to insert a register block in the middle of another sequential file:

```
struct debugfs_reg32 {
    char *name;
    unsigned long offset;
};

struct debugfs_regset32 {
    const struct debugfs_reg32 *regs;
    int nregs;
    void __iomem *base;
    struct device *dev;     /* Optional device for Runtime PM */
};

debugfs_create_regset32(const char *name, umode_t mode,
                        struct dentry *parent,
                        struct debugfs_regset32 *regset);

void debugfs_print_regs32(struct seq_file *s, const struct debugfs_reg32 *regs,
                      int nregs, void __iomem *base, char *prefix);
```

The ”base” argument may be 0, but you may want to build the reg32 array using __stringify, and a number of register names (macros) are actually byte offsets over a base for the register block.

If you want to dump a u32 array in debugfs, you can create a file with:

```
struct debugfs_u32_array {
    u32 *array;
    u32 n_elements;
};

void debugfs_create_u32_array(const char *name, umode_t mode,
                    struct dentry *parent,
                    struct debugfs_u32_array *array);
```

The ”array” argument wraps a pointer to the array's data and the number of its elements. Note: Once array is created its size can not be changed.

There is a helper function to create a device-related seq_file:

```
void debugfs_create_devm_seqfile(struct device *dev,
                                 const char *name,
                                 struct dentry *parent,
                                 int (*read_fn)(struct seq_file *s,
                                         void *data));
```

The ”dev” argument is the device related to this debugfs file, and the ”read_fn” is a function pointer which to be called to print the seq_file content.

There are a couple of other directory-oriented helper functions:

```
struct dentry *debugfs_rename(struct dentry *old_dir,
                              struct dentry *old_dentry,
                              struct dentry *new_dir,
                              const char *new_name);

struct dentry *debugfs_create_symlink(const char *name,
                                      struct dentry *parent,
                                      const char *target);
```

A call to *debugfs_rename()* will give a new name to an existing debugfs file, possibly in a different directory. The new_name must not exist prior to the call; the return value is old_dentry with updated information. Symbolic links can be created with *debugfs_create_symlink()*.

There is one important thing that all debugfs users must take into account: there is no automatic cleanup of any directories created in debugfs. If a module is unloaded without explicitly removing debugfs entries, the result will be a lot of stale pointers and no end of highly antisocial behavior. So all debugfs users - at least those which can be built as modules - must be prepared to remove all files and directories they create there. A file can be removed with:

```
void debugfs_remove(struct dentry *dentry);
```

The dentry value can be NULL or an error value, in which case nothing will be removed.

Once upon a time, debugfs users were required to remember the dentry pointer for every debugfs file they created so that all files could be cleaned up. We live in more civilized times now, though, and debugfs users can call:

```
void debugfs_remove_recursive(struct dentry *dentry);
```

If this function is passed a pointer for the dentry corresponding to the top-level directory, the entire hierarchy below that directory will be removed.

## 3.16 DLMFS

A minimal DLM userspace interface implemented via a virtual file system.

dlmfs is built with OCFS2 as it requires most of its infrastructure.

**Project web page**
http://ocfs2.wiki.kernel.org

**Tools web page**
https://github.com/markfasheh/ocfs2-tools

**OCFS2 mailing lists**
https://subspace.kernel.org/lists.linux.dev.html



### 3.16.1 Credits

Some code taken from ramfs which is Copyright © 2000 Linus Torvalds and Transmeta Corp.

Mark Fasheh <mark.fasheh@oracle.com>

### 3.16.2 Caveats

- Right now it only works with the OCFS2 DLM, though support for other DLM implementations should not be a major issue.

### 3.16.3 Mount options

None

### 3.16.4 Usage

If you're just interested in OCFS2, then please see *OCFS2 filesystem*. The rest of this document will be geared towards those who want to use dlmfs for easy to setup and easy to use clustered locking in userspace.

### 3.16.5 Setup

dlmfs requires that the OCFS2 cluster infrastructure be in place. Please download ocfs2-tools from the above url and configure a cluster.

You'll want to start heartbeating on a volume which all the nodes in your lockspace can access. The easiest way to do this is via ocfs2_hb_ctl (distributed with ocfs2-tools). Right now it requires that an OCFS2 file system be in place so that it can automatically find its heartbeat area, though it will eventually support heartbeat against raw disks.

Please see the ocfs2_hb_ctl and mkfs.ocfs2 manual pages distributed with ocfs2-tools.

Once you're heartbeating, DLM lock 'domains' can be easily created / destroyed and locks within them accessed.

### 3.16.6 Locking

Users may access dlmfs via standard file system calls, or they can use 'libo2dlm' (distributed with ocfs2-tools) which abstracts the file system calls and presents a more traditional locking api.

dlmfs handles lock caching automatically for the user, so a lock request for an already acquired lock will not generate another DLM call. Userspace programs are assumed to handle their own local locking.

Two levels of locks are supported - Shared Read, and Exclusive. Also supported is a Trylock operation.

For information on the libo2dlm interface, please see o2dlm.h, distributed with ocfs2-tools.

Lock value blocks can be read and written to a resource via read(2) and write(2) against the fd obtained via your open(2) call. The maximum currently supported LVB length is 64 bytes (though that is an OCFS2 DLM limitation). Through this mechanism, users of dlmfs can share small amounts of data amongst their nodes.

mkdir(2) signals dlmfs to join a domain (which will have the same name as the resulting directory)

rmdir(2) signals dlmfs to leave the domain

Locks for a given domain are represented by regular inodes inside the domain directory. Locking against them is done via the open(2) system call.

The open(2) call will not return until your lock has been granted or an error has occurred, unless it has been instructed to do a trylock operation. If the lock succeeds, you'll get an fd.

open(2) with O_CREAT to ensure the resource inode is created - dlmfs does not automatically create inodes for existing lock resources.

| Open Flag | Lock Request Type |
|---|---|
| O_RDONLY | Shared Read |
| O_RDWR | Exclusive |

| Open Flag | Resulting Locking Behavior |
|---|---|
| O_NONBLOCK | Trylock operation |

You must provide exactly one of O_RDONLY or O_RDWR.

If O_NONBLOCK is also provided and the trylock operation was valid but could not lock the resource then open(2) will return ETXTBUSY.

close(2) drops the lock associated with your fd.

Modes passed to mkdir(2) or open(2) are adhered to locally. Chown is supported locally as well. This means you can use them to restrict access to the resources via dlmfs on your local node only.

The resource LVB may be read from the fd in either Shared Read or Exclusive modes via the read(2) system call. It can be written via write(2) only when open in Exclusive mode.

Once written, an LVB will be visible to other nodes who obtain Read Only or higher level locks on the resource.

### 3.16.7 See Also

http://opendlm.sourceforge.net/cvsmirror/opendlm/docs/dlmbook_final.pdf

For more information on the VMS distributed locking API.

## 3.17 eCryptfs: A stacked cryptographic filesystem for Linux

eCryptfs is free software. Please see the file COPYING for details. For documentation, please see the files in the doc/ subdirectory. For building and installation instructions please see the INSTALL file.

**Maintainer**
Phillip Hellewell

**Lead developer**
Michael A. Halcrow <mhalcrow@us.ibm.com>

**Developers**
Michael C. Thompson Kent Yoder

**Web Site**
http://ecryptfs.sf.net

This software is currently undergoing development. Make sure to maintain a backup copy of any data you write into eCryptfs.

eCryptfs requires the userspace tools downloadable from the SourceForge site:

http://sourceforge.net/projects/ecryptfs/

Userspace requirements include:

- David Howells' userspace keyring headers and libraries (version 1.0 or higher), obtainable from http://people.redhat.com/~dhowells/keyutils/
- Libgcrypt

**Note:** In the beta/experimental releases of eCryptfs, when you upgrade eCryptfs, you should copy the files to an unencrypted location and then copy the files back into the new eCryptfs mount to migrate the files.

### 3.17.1 Mount-wide Passphrase

Create a new directory into which eCryptfs will write its encrypted files (i.e., /root/crypt). Then, create the mount point directory (i.e., /mnt/crypt). Now it's time to mount eCryptfs:

```
mount -t ecryptfs /root/crypt /mnt/crypt
```

You should be prompted for a passphrase and a salt (the salt may be blank).

Try writing a new file:

```
echo "Hello, World" > /mnt/crypt/hello.txt
```

The operation will complete. Notice that there is a new file in /root/crypt that is at least 12288 bytes in size (depending on your host page size). This is the encrypted underlying file for what you just wrote. To test reading, from start to finish, you need to clear the user session keyring:

keyctl clear @u

Then umount /mnt/crypt and mount again per the instructions given above.

```
cat /mnt/crypt/hello.txt
```

### 3.17.2 Notes

eCryptfs version 0.1 should only be mounted on (1) empty directories or (2) directories containing files only created by eCryptfs. If you mount a directory that has pre-existing files not created by eCryptfs, then behavior is undefined. Do not run eCryptfs in higher verbosity levels unless you are doing so for the sole purpose of debugging or development, since secret values will be written out to the system log in that case.

Mike Halcrow mhalcrow@us.ibm.com

## 3.18 efivarfs - a (U)EFI variable filesystem

The efivarfs filesystem was created to address the shortcomings of using entries in sysfs to maintain EFI variables. The old sysfs EFI variables code only supported variables of up to 1024 bytes. This limitation existed in version 0.99 of the EFI specification, but was removed before any full releases. Since variables can now be larger than a single page, sysfs isn't the best interface for this.

Variables can be created, deleted and modified with the efivarfs filesystem.

efivarfs is typically mounted like this:

```
mount -t efivarfs none /sys/firmware/efi/efivars
```

Due to the presence of numerous firmware bugs where removing non-standard UEFI variables causes the system firmware to fail to POST, efivarfs files that are not well-known standardized variables are created as immutable files. This doesn't prevent removal - ”chattr -i” will work - but it does prevent this kind of failure from being accomplished accidentally.

**Warning:** When a content of an UEFI variable in /sys/firmware/efi/efivars is displayed, for example using ”hexdump”, pay attention that the first 4 bytes of the output represent the UEFI variable attributes, in little-endian format.

Practically the output of each efivar is composed of:

```
4_bytes_of_attributes + efivar_data
```

*See also:*

- Documentation/admin-guide/acpi/ssdt-overlays.rst
- Documentation/ABI/stable/sysfs-firmware-efi-vars

## 3.19 EROFS - Enhanced Read-Only File System

### 3.19.1 Overview

EROFS filesystem stands for Enhanced Read-Only File System. It aims to form a generic read-only filesystem solution for various read-only use cases instead of just focusing on storage space saving without considering any side effects of runtime performance.

It is designed to meet the needs of flexibility, feature extendability and user payload friendly, etc. Apart from those, it is still kept as a simple random-access friendly high-performance filesystem to get rid of unneeded I/O amplification and memory-resident overhead compared to similar approaches.

It is implemented to be a better choice for the following scenarios:

- read-only storage media or
- part of a fully trusted read-only solution, which means it needs to be immutable and bit-for-bit identical to the official golden image for their releases due to security or other considerations and
- hope to minimize extra storage space with guaranteed end-to-end performance by using compact layout, transparent file compression and direct access, especially for those embedded devices with limited memory and high-density hosts with numerous containers.

Here are the main features of EROFS:

- Little endian on-disk design;
- Block-based distribution and file-based distribution over fscache are supported;
- Support multiple devices to refer to external blobs, which can be used for container images;

- 32-bit block addresses for each device, therefore 16TiB address space at most with 4KiB block size for now;
- Two inode layouts for different requirements:

| | | |
|---|---|---|
| Inode metadata size | 32 bytes | 64 bytes |
| Max file size | 4 GiB | 16 EiB (also limited by max. vol size) |
| Max uids/gids | 65536 | 4294967296 |
| Per-inode timestamp | no | yes (64 + 32-bit timestamp) |
| Max hardlinks | 65536 | 4294967296 |
| Metadata reserved | 8 bytes | 18 bytes |

- Support extended attributes as an option;
- Support a bloom filter that speeds up negative extended attribute lookups;
- Support POSIX.1e ACLs by using extended attributes;
- Support transparent data compression as an option: LZ4, MicroLZMA and DEFLATE algorithms can be used on a per-file basis; In addition, inplace decompression is also supported to avoid bounce compressed buffers and unnecessary page cache thrashing.
- Support chunk-based data deduplication and rolling-hash compressed data deduplication;
- Support tailpacking inline compared to byte-addressed unaligned metadata or smaller block size alternatives;
- Support merging tail-end data into a special inode as fragments.
- Support large folios for uncompressed files.
- Support direct I/O on uncompressed files to avoid double caching for loop devices;
- Support FSDAX on uncompressed images for secure containers and ramdisks in order to get rid of unnecessary page cache.
- Support file-based on-demand loading with the Fscache infrastructure.

The following git tree provides the file system user-space tools under development, such as a formatting tool (mkfs.erofs), an on-disk consistency & compatibility checking tool (fsck.erofs), and a debugging tool (dump.erofs):

- git://git.kernel.org/pub/scm/linux/kernel/git/xiang/erofs-utils.git

For more information, please also refer to the documentation site:

- https://erofs.docs.kernel.org

Bugs and patches are welcome, please kindly help us and send to the following linux-erofs mailing list:

- linux-erofs mailing list <linux-erofs@lists.ozlabs.org>

### 3.19.2 Mount options

<table>
<tr><td>(no)user_xattr</td><td>Setup Extended User Attributes. Note: xattr is enabled by default if CONFIG_EROFS_FS_XATTR is selected.</td></tr>
<tr><td>(no)acl</td><td>Setup POSIX Access Control List. Note: acl is enabled by default if CONFIG_EROFS_FS_POSIX_ACL is selected.</td></tr>
<tr><td>cache_strategy=%s</td><td>Select a strategy for cached decompression from now on:<br><br><table><tr><td>disabled</td><td>In-place I/O decompression only;</td></tr><tr><td>readahead</td><td>Cache the last incomplete compressed physical cluster for further reading. It still does in-place I/O decompression for the rest compressed physical clusters;</td></tr><tr><td>readaround</td><td>Cache the both ends of incomplete compressed physical clusters for further reading. It still does in-place I/O decompression for the rest compressed physical clusters.</td></tr></table></td></tr>
<tr><td>dax={always,never}</td><td>Use direct access (no page cache). See Direct Access for files.</td></tr>
<tr><td>dax</td><td>A legacy option which is an alias for <code>dax=always</code>.</td></tr>
<tr><td>device=%s</td><td>Specify a path to an extra device to be used together.</td></tr>
<tr><td>fsid=%s</td><td>Specify a filesystem image ID for Fscache back-end.</td></tr>
<tr><td>domain_id=%s</td><td>Specify a domain ID in fscache mode so that different images with the same blobs under a given domain ID can share storage.</td></tr>
</table>

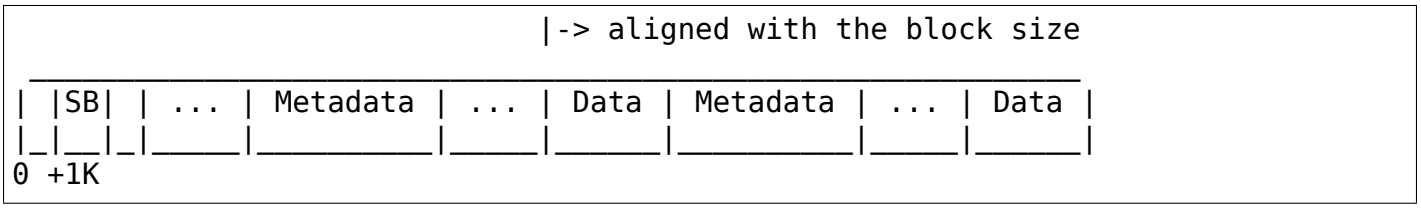
### 3.19.3 Sysfs Entries

Information about mounted erofs file systems can be found in /sys/fs/erofs. Each mounted filesystem will have a directory in /sys/fs/erofs based on its device name (i.e., /sys/fs/erofs/sda). (see also Documentation/ABI/testing/sysfs-fs-erofs)

### 3.19.4 On-disk details

#### Summary

Different from other read-only file systems, an EROFS volume is designed to be as simple as possible:

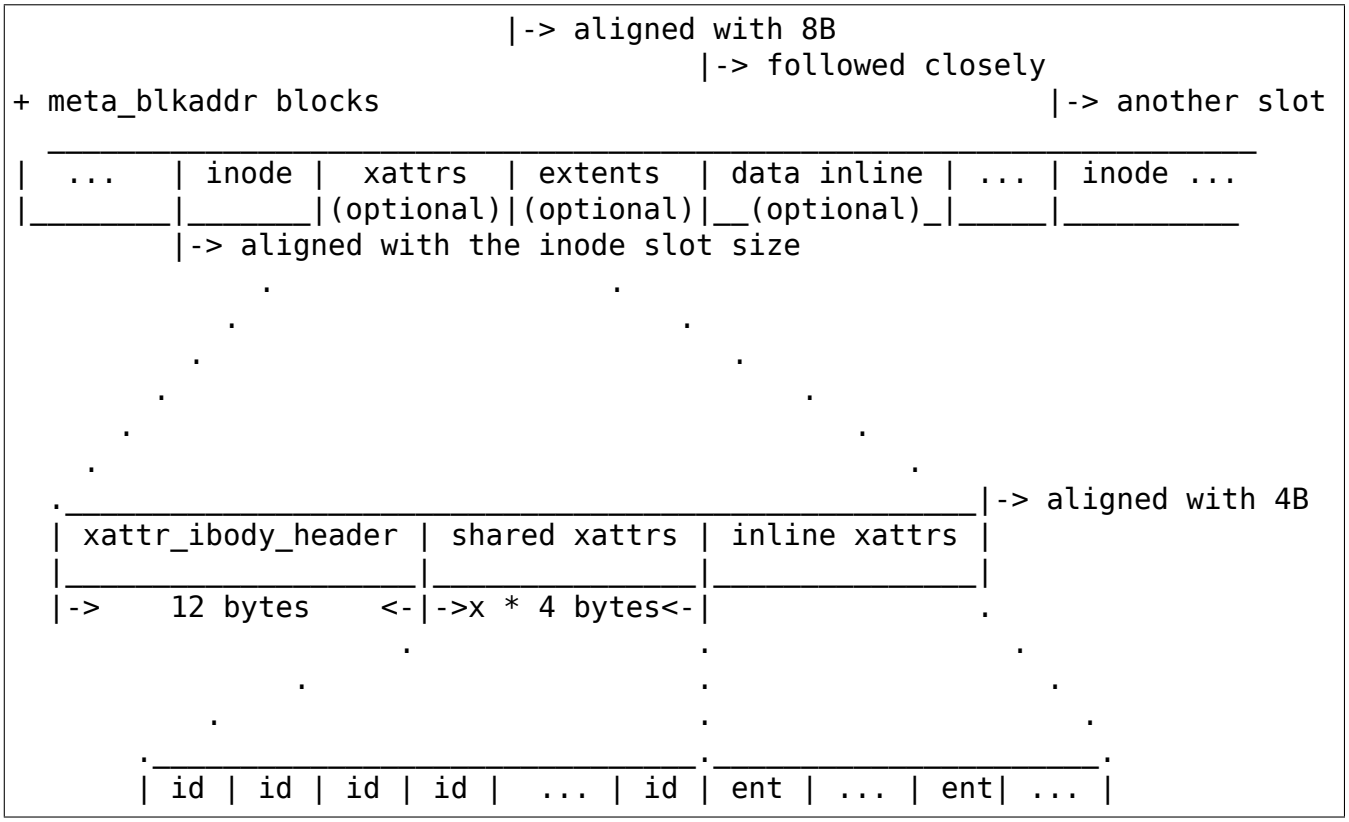


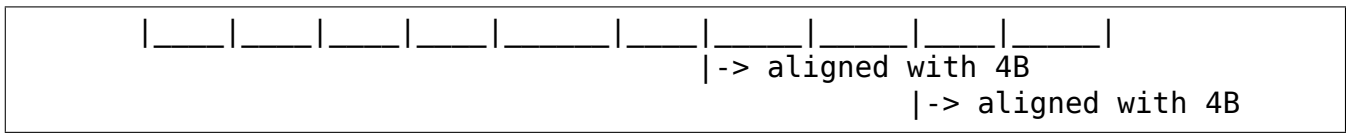

All data areas should be aligned with the block size, but metadata areas may not. All metadatas can be now observed in two different spaces (views):

1. Inode metadata space

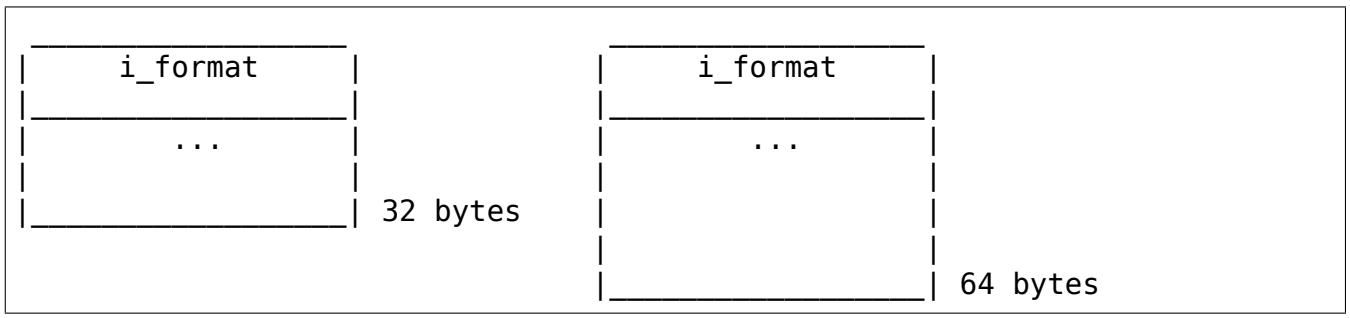
   Each valid inode should be aligned with an inode slot, which is a fixed value (32 bytes) and designed to be kept in line with compact inode size.

   **Each inode can be directly found with the following formula:**
   inode offset = meta_blkaddr * block_size + 32 * nid






Inode could be 32 or 64 bytes, which can be distinguished from a common field which all inode versions have -- i_format:




Xattrs, extents, data inline are placed after the corresponding inode with proper alignment, and they could be optional for different data mappings. _currently_ total 5 data layouts are supported:

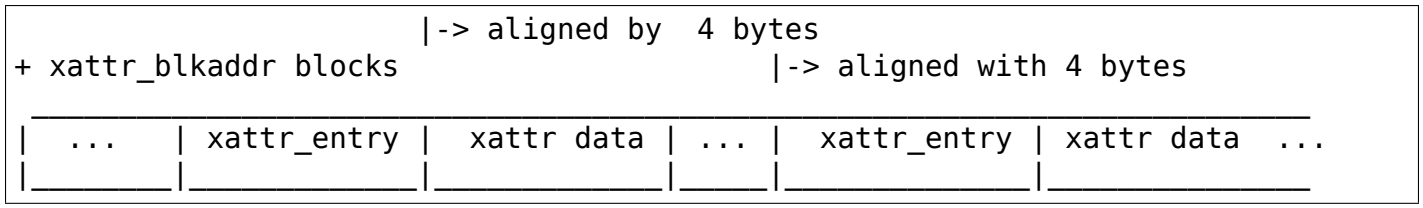
| | |
|---|---|
| 0 | flat file data without data inline (no extent); |
| 1 | fixed-sized output data compression (with non-compacted indexes); |
| 2 | flat file data with tail packing data inline (no extent); |
| 3 | fixed-sized output data compression (with compacted indexes, v5.3+); |
| 4 | chunk-based file (v5.15+). |

The size of the optional xattrs is indicated by i_xattr_count in inode header. Large xattrs or xattrs shared by many different files can be stored in shared xattrs metadata rather than inlined right after inode.

2. Shared xattrs metadata space

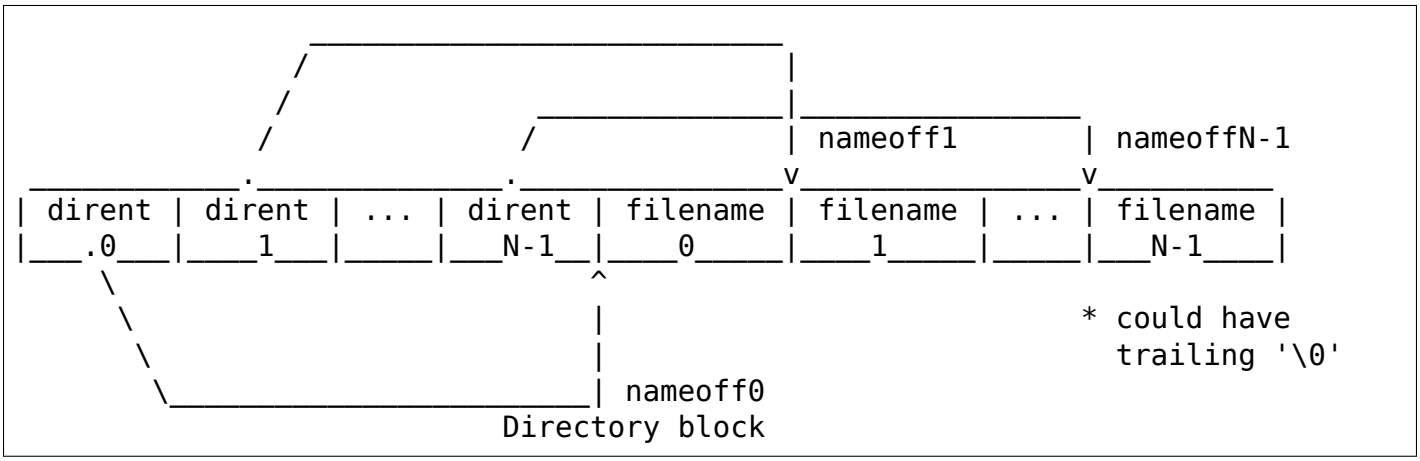
   Shared xattrs space is similar to the above inode space, started with a specific block indicated by xattr_blkaddr, organized one by one with proper align.

   **Each share xattr can also be directly found by the following formula:**
   xattr offset = xattr_blkaddr * block_size + 4 * xattr_id



### Directories

All directories are now organized in a compact on-disk format. Note that each directory block is divided into index and name areas in order to support random file lookup, and all directory entries are _strictly_ recorded in alphabetical order in order to support improved prefix binary search algorithm (could refer to the related source code).




Note that apart from the offset of the first filename, nameoff0 also indicates the total number of directory entries in this block since it is no need to introduce another on-disk field at all.

### Chunk-based files

In order to support chunk-based data deduplication, a new inode data layout has been supported since Linux v5.15: Files are split in equal-sized data chunks with `extents` area of the inode metadata indicating how to get the chunk data: these can be simply as a 4-byte block address array or in the 8-byte chunk index form (see struct erofs_inode_chunk_index in erofs_fs.h for more details.)

By the way, chunk-based files are all uncompressed for now.

### Long extended attribute name prefixes

There are use cases where extended attributes with different values can have only a few common prefixes (such as overlayfs xattrs). The predefined prefixes work inefficiently in both image size and runtime performance in such cases.

The long xattr name prefixes feature is introduced to address this issue. The overall idea is that, apart from the existing predefined prefixes, the xattr entry could also refer to user-specified long xattr name prefixes, e.g. ”trusted.overlay.”.

When referring to a long xattr name prefix, the highest bit (bit 7) of erofs_xattr_entry.e_name_index is set, while the lower bits (bit 0-6) as a whole represent the index of the referred long name prefix among all long name prefixes. Therefore, only the trailing part of the name apart from the long xattr name prefix is stored in erofs_xattr_entry.e_name, which could be empty if the full xattr name matches exactly as its long xattr name prefix.

All long xattr prefixes are stored one by one in the packed inode as long as the packed inode is valid, or in the meta inode otherwise. The xattr_prefix_count (of the on-disk superblock) indicates the total number of long xattr name prefixes, while (xattr_prefix_start * 4) indicates the

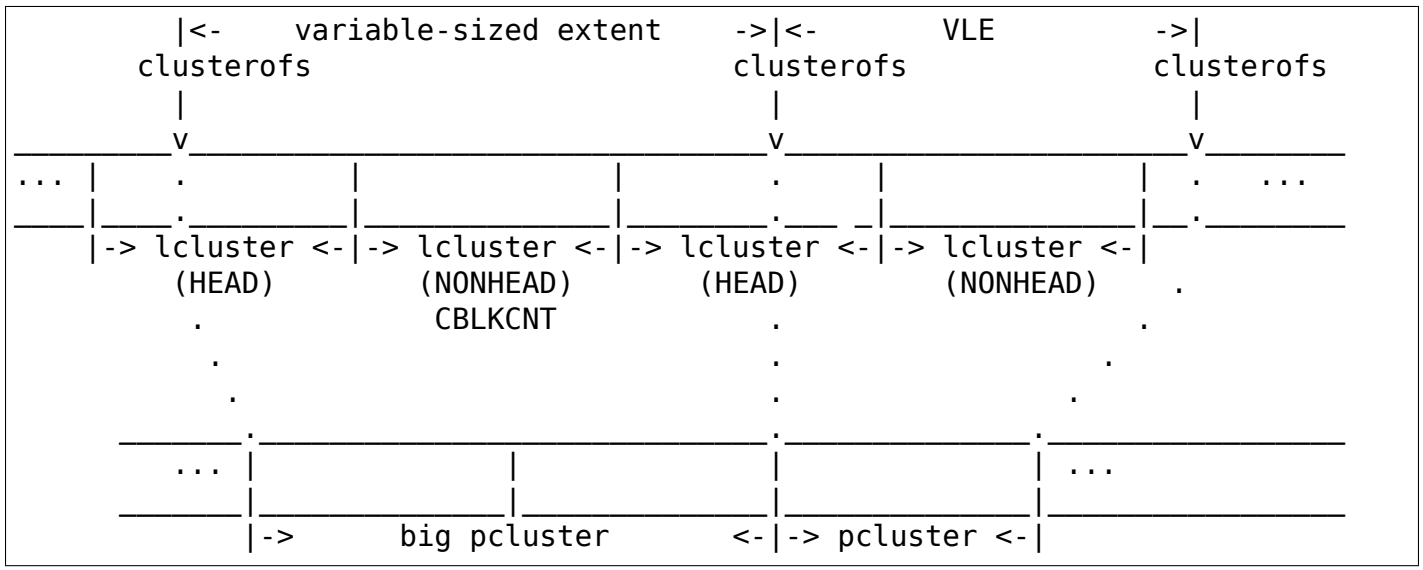
start offset of long name prefixes in the packed/meta inode. Note that, long extended attribute name prefixes are disabled if xattr_prefix_count is 0.

Each long name prefix is stored in the format: ALIGN({__le16 len, data}, 4), where len represents the total size of the data part. The data part is actually represented by 'struct erofs_xattr_long_prefix', where base_index represents the index of the predefined xattr name prefix, e.g. EROFS_XATTR_INDEX_TRUSTED for ”trusted.overlay.” long name prefix, while the infix string keeps the string after stripping the short prefix, e.g. ”overlay.” for the example above.

### Data compression

EROFS implements fixed-sized output compression which generates fixed-sized compressed data blocks from variable-sized input in contrast to other existing fixed-sized input solutions. Relatively higher compression ratios can be gotten by using fixed-sized output compression since nowadays popular data compression algorithms are mostly LZ77-based and such fixed-sized output approach can be benefited from the historical dictionary (aka. sliding window).

In details, original (uncompressed) data is turned into several variable-sized extents and in the meanwhile, compressed into physical clusters (pclusters). In order to record each variable-sized extent, logical clusters (lclusters) are introduced as the basic unit of compress indexes to indicate whether a new extent is generated within the range (HEAD) or not (NONHEAD). Lclusters are now fixed in block size, as illustrated below:

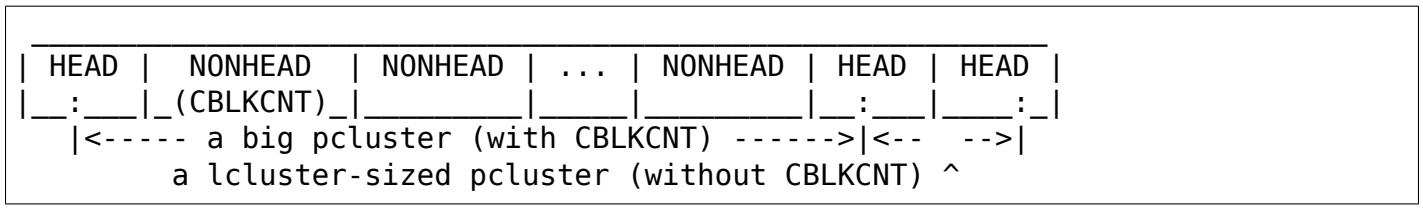



A physical cluster can be seen as a container of physical compressed blocks which contains compressed data. Previously, only lcluster-sized (4KB) pclusters were supported. After big pcluster feature is introduced (available since Linux v5.13), pcluster can be a multiple of lcluster size.

For each HEAD lcluster, clusterofs is recorded to indicate where a new extent starts and blkaddr is used to seek the compressed data. For each NONHEAD lcluster, delta0 and delta1 are available instead of blkaddr to indicate the distance to its HEAD lcluster and the next HEAD lcluster. A PLAIN lcluster is also a HEAD lcluster except that its data is uncompressed. See the comments around ”struct z_erofs_vle_decompressed_index” in erofs_fs.h for more details.

If big pcluster is enabled, pcluster size in lclusters needs to be recorded as well. Let the delta0 of the first NONHEAD lcluster store the compressed block count with a special flag as a new called

CBLKCNT NONHEAD lcluster. It's easy to understand its delta0 is constantly 1, as illustrated below:




If another HEAD follows a HEAD lcluster, there is no room to record CBLKCNT, but it's easy to know the size of such pcluster is 1 lcluster as well.

Since Linux v6.1, each pcluster can be used for multiple variable-sized extents, therefore it can be used for compressed data deduplication.

## 3.20 The Second Extended Filesystem

ext2 was originally released in January 1993. Written by R'emy Card, Theodore Ts'o and Stephen Tweedie, it was a major rewrite of the Extended Filesystem. It is currently still (April 2001) the predominant filesystem in use by Linux. There are also implementations available for NetBSD, FreeBSD, the GNU HURD, Windows 95/98/NT, OS/2 and RISC OS.

### 3.20.1 Options

Most defaults are determined by the filesystem superblock, and can be set using tune2fs(8). Kernel-determined defaults are indicated by (*).

| | | |
|---|---|---|
| bsddf | (*) | Makes `df` act like BSD. |
| minixdf | | Makes `df` act like Minix. |
| check=none, nocheck | (*) | Don't do extra checking of bitmaps on mount (check=normal and check=strict options removed) |
| dax | | Use direct access (no page cache). See *Direct Access for files*. |
| debug | | Extra debugging information is sent to the kernel syslog. Useful for developers. |
| errors=continue | | Keep going on a filesystem error. |
| errors=remount-ro | | Remount the filesystem read-only on an error. |
| errors=panic | | Panic and halt the machine if an error occurs. |
| grpid, bsdgroups | | Give objects the same group ID as their parent. |
| nogrpid, sysvgroups | | New objects have the group ID of their creator. |
| nouid32 | | Use 16-bit UIDs and GIDs. |
| oldalloc | | Enable the old block allocator. Orlov should have better performance, we'd like to get some feedback if it's the contrary for you. |
| orlov | (*) | Use the Orlov block allocator. (See http://lwn.net/Articles/14633/ and http://lwn.net/Articles/14446/.) |
| resuid=n | | The user ID which may use the reserved blocks. |
| resgid=n | | The group ID which may use the reserved blocks. |
| sb=n | | Use alternate superblock at this location. |
| user_xattr | | Enable "user." POSIX Extended Attributes (requires CONFIG_EXT2_FS_XATTR). |
| nouser_xattr | | Don't support "user." extended attributes. |
| acl | | Enable POSIX Access Control Lists support (requires CONFIG_EXT2_FS_POSIX_ACL). |
| noacl | | Don't support POSIX ACLs. |
| quota, usrquota | | Enable user disk quota support (requires CONFIG_QUOTA). |
| grpquota | | Enable group disk quota support (requires CONFIG_QUOTA). |

noquota option ls silently ignored by ext2.

### 3.20.2 Specification

ext2 shares many properties with traditional Unix filesystems. It has the concepts of blocks, inodes and directories. It has space in the specification for Access Control Lists (ACLs), fragments, undeletion and compression though these are not yet implemented (some are available as separate patches). There is also a versioning mechanism to allow new features (such as journalling) to be added in a maximally compatible manner.

### Blocks

The space in the device or file is split up into blocks. These are a fixed size, of 1024, 2048 or 4096 bytes (8192 bytes on Alpha systems), which is decided when the filesystem is created. Smaller blocks mean less wasted space per file, but require slightly more accounting overhead, and also impose other limits on the size of files and the filesystem.

### Block Groups

Blocks are clustered into block groups in order to reduce fragmentation and minimise the amount of head seeking when reading a large amount of consecutive data. Information about each block group is kept in a descriptor table stored in the block(s) immediately after the superblock. Two blocks near the start of each group are reserved for the block usage bitmap and the inode usage bitmap which show which blocks and inodes are in use. Since each bitmap is limited to a single block, this means that the maximum size of a block group is 8 times the size of a block.

The block(s) following the bitmaps in each block group are designated as the inode table for that block group and the remainder are the data blocks. The block allocation algorithm attempts to allocate data blocks in the same block group as the inode which contains them.

### The Superblock

The superblock contains all the information about the configuration of the filing system. The primary copy of the superblock is stored at an offset of 1024 bytes from the start of the device, and it is essential to mounting the filesystem. Since it is so important, backup copies of the superblock are stored in block groups throughout the filesystem. The first version of ext2 (revision 0) stores a copy at the start of every block group, along with backups of the group descriptor block(s). Because this can consume a considerable amount of space for large filesystems, later revisions can optionally reduce the number of backup copies by only putting backups in specific groups (this is the sparse superblock feature). The groups chosen are 0, 1 and powers of 3, 5 and 7.

The information in the superblock contains fields such as the total number of inodes and blocks in the filesystem and how many are free, how many inodes and blocks are in each block group, when the filesystem was mounted (and if it was cleanly unmounted), when it was modified, what version of the filesystem it is (see the Revisions section below) and which OS created it.

If the filesystem is revision 1 or higher, then there are extra fields, such as a volume name, a unique identification number, the inode size, and space for optional filesystem features to store configuration info.

All fields in the superblock (as in all other ext2 structures) are stored on the disc in little endian format, so a filesystem is portable between machines without having to know what machine it was created on.

### Inodes

The inode (index node) is a fundamental concept in the ext2 filesystem. Each object in the filesystem is represented by an inode. The inode structure contains pointers to the filesystem blocks which contain the data held in the object and all of the metadata about an object except its name. The metadata about an object includes the permissions, owner, group, flags, size, number of blocks used, access time, change time, modification time, deletion time, number of links, fragments, version (for NFS) and extended attributes (EAs) and/or Access Control Lists (ACLs).

There are some reserved fields which are currently unused in the inode structure and several which are overloaded. One field is reserved for the directory ACL if the inode is a directory and alternately for the top 32 bits of the file size if the inode is a regular file (allowing file sizes larger than 2GB). The translator field is unused under Linux, but is used by the HURD to reference the inode of a program which will be used to interpret this object. Most of the remaining reserved fields have been used up for both Linux and the HURD for larger owner and group fields, The HURD also has a larger mode field so it uses another of the remaining fields to store the extra more bits.

There are pointers to the first 12 blocks which contain the file's data in the inode. There is a pointer to an indirect block (which contains pointers to the next set of blocks), a pointer to a doubly-indirect block (which contains pointers to indirect blocks) and a pointer to a trebly-indirect block (which contains pointers to doubly-indirect blocks).

The flags field contains some ext2-specific flags which aren't catered for by the standard chmod flags. These flags can be listed with lsattr and changed with the chattr command, and allow specific filesystem behaviour on a per-file basis. There are flags for secure deletion, undeletable, compression, synchronous updates, immutability, append-only, dumpable, no-atime, indexed directories, and data-journaling. Not all of these are supported yet.

### Directories

A directory is a filesystem object and has an inode just like a file. It is a specially formatted file containing records which associate each name with an inode number. Later revisions of the filesystem also encode the type of the object (file, directory, symlink, device, fifo, socket) to avoid the need to check the inode itself for this information (support for taking advantage of this feature does not yet exist in Glibc 2.2).

The inode allocation code tries to assign inodes which are in the same block group as the directory in which they are first created.

The current implementation of ext2 uses a singly-linked list to store the filenames in the directory; a pending enhancement uses hashing of the filenames to allow lookup without the need to scan the entire directory.

The current implementation never removes empty directory blocks once they have been allocated to hold more files.

### Special files

Symbolic links are also filesystem objects with inodes. They deserve special mention because the data for them is stored within the inode itself if the symlink is less than 60 bytes long. It uses the fields which would normally be used to store the pointers to data blocks. This is a worthwhile optimisation as it we avoid allocating a full block for the symlink, and most symlinks are less than 60 characters long.

Character and block special devices never have data blocks assigned to them. Instead, their device number is stored in the inode, again reusing the fields which would be used to point to the data blocks.

### Reserved Space

In ext2, there is a mechanism for reserving a certain number of blocks for a particular user (normally the super-user). This is intended to allow for the system to continue functioning even if non-privileged users fill up all the space available to them (this is independent of filesystem quotas). It also keeps the filesystem from filling up entirely which helps combat fragmentation.

### Filesystem check

At boot time, most systems run a consistency check (e2fsck) on their filesystems. The superblock of the ext2 filesystem contains several fields which indicate whether fsck should actually run (since checking the filesystem at boot can take a long time if it is large). fsck will run if the filesystem was not cleanly unmounted, if the maximum mount count has been exceeded or if the maximum time between checks has been exceeded.

### Feature Compatibility

The compatibility feature mechanism used in ext2 is sophisticated. It safely allows features to be added to the filesystem, without unnecessarily sacrificing compatibility with older versions of the filesystem code. The feature compatibility mechanism is not supported by the original revision 0 (EXT2_GOOD_OLD_REV) of ext2, but was introduced in revision 1. There are three 32-bit fields, one for compatible features (COMPAT), one for read-only compatible (RO_COMPAT) features and one for incompatible (INCOMPAT) features.

These feature flags have specific meanings for the kernel as follows:

A COMPAT flag indicates that a feature is present in the filesystem, but the on-disk format is 100% compatible with older on-disk formats, so a kernel which didn't know anything about this feature could read/write the filesystem without any chance of corrupting the filesystem (or even making it inconsistent). This is essentially just a flag which says ”this filesystem has a (hidden) feature” that the kernel or e2fsck may want to be aware of (more on e2fsck and feature flags later). The ext3 HAS_JOURNAL feature is a COMPAT flag because the ext3 journal is simply a regular file with data blocks in it so the kernel does not need to take any special notice of it if it doesn't understand ext3 journaling.

An RO_COMPAT flag indicates that the on-disk format is 100% compatible with older on-disk formats for reading (i.e. the feature does not change the visible on-disk format). However, an old kernel writing to such a filesystem would/could corrupt the filesystem, so this is prevented. The most common such feature, SPARSE_SUPER, is an RO_COMPAT feature because sparse

groups allow file data blocks where superblock/group descriptor backups used to live, and ext2_free_blocks() refuses to free these blocks, which would leading to inconsistent bitmaps. An old kernel would also get an error if it tried to free a series of blocks which crossed a group boundary, but this is a legitimate layout in a SPARSE_SUPER filesystem.

An INCOMPAT flag indicates the on-disk format has changed in some way that makes it unreadable by older kernels, or would otherwise cause a problem if an old kernel tried to mount it. FILETYPE is an INCOMPAT flag because older kernels would think a filename was longer than 256 characters, which would lead to corrupt directory listings. The COMPRESSION flag is an obvious INCOMPAT flag - if the kernel doesn't understand compression, you would just get garbage back from read() instead of it automatically decompressing your data. The ext3 RECOVER flag is needed to prevent a kernel which does not understand the ext3 journal from mounting the filesystem without replaying the journal.

For e2fsck, it needs to be more strict with the handling of these flags than the kernel. If it doesn't understand ANY of the COMPAT, RO_COMPAT, or INCOMPAT flags it will refuse to check the filesystem, because it has no way of verifying whether a given feature is valid or not. Allowing e2fsck to succeed on a filesystem with an unknown feature is a false sense of security for the user. Refusing to check a filesystem with unknown features is a good incentive for the user to update to the latest e2fsck. This also means that anyone adding feature flags to ext2 also needs to update e2fsck to verify these features.

### Metadata

It is frequently claimed that the ext2 implementation of writing asynchronous metadata is faster than the ffs synchronous metadata scheme but less reliable. Both methods are equally resolvable by their respective fsck programs.

If you're exceptionally paranoid, there are 3 ways of making metadata writes synchronous on ext2:

- per-file if you have the program source: use the O_SYNC flag to open()
- per-file if you don't have the source: use ”chattr +S” on the file
- per-filesystem: add the ”sync” option to mount (or in /etc/fstab)

the first and last are not ext2 specific but do force the metadata to be written synchronously. See also Journaling below.

### Limitations

There are various limits imposed by the on-disk layout of ext2. Other limits are imposed by the current implementation of the kernel code. Many of the limits are determined at the time the filesystem is first created, and depend upon the block size chosen. The ratio of inodes to data blocks is fixed at filesystem creation time, so the only way to increase the number of inodes is to increase the size of the filesystem. No tools currently exist which can change the ratio of inodes to blocks.

Most of these limits could be overcome with slight changes in the on-disk format and using a compatibility flag to signal the format change (at the expense of some compatibility).

| Filesystem block size | 1kB | 2kB | 4kB | 8kB |
|---|---|---|---|---|
| File size limit | 16GB | 256GB | 2048GB | 2048GB |
| Filesystem size limit | 2047GB | 8192GB | 16384GB | 32768GB |

There is a 2.4 kernel limit of 2048GB for a single block device, so no filesystem larger than that can be created at this time. There is also an upper limit on the block size imposed by the page size of the kernel, so 8kB blocks are only allowed on Alpha systems (and other architectures which support larger pages).

There is an upper limit of 32000 subdirectories in a single directory.

There is a ”soft” upper limit of about 10-15k files in a single directory with the current linear linked-list directory implementation. This limit stems from performance problems when creating and deleting (and also finding) files in such large directories. Using a hashed directory index (under development) allows 100k-1M+ files in a single directory without performance problems (although RAM size becomes an issue at this point).

The (meaningless) absolute upper limit of files in a single directory (imposed by the file size, the realistic limit is obviously much less) is over 130 trillion files. It would be higher except there are not enough 4-character names to make up unique directory entries, so they have to be 8 character filenames, even then we are fairly close to running out of unique filenames.

#### Journaling

A journaling extension to the ext2 code has been developed by Stephen Tweedie. It avoids the risks of metadata corruption and the need to wait for e2fsck to complete after a crash, without requiring a change to the on-disk ext2 layout. In a nutshell, the journal is a regular file which stores whole metadata (and optionally data) blocks that have been modified, prior to writing them into the filesystem. This means it is possible to add a journal to an existing ext2 filesystem without the need for data conversion.

When changes to the filesystem (e.g. a file is renamed) they are stored in a transaction in the journal and can either be complete or incomplete at the time of a crash. If a transaction is complete at the time of a crash (or in the normal case where the system does not crash), then any blocks in that transaction are guaranteed to represent a valid filesystem state, and are copied into the filesystem. If a transaction is incomplete at the time of the crash, then there is no guarantee of consistency for the blocks in that transaction so they are discarded (which means any filesystem changes they represent are also lost). Check Documentation/filesystems/ext4/ if you want to read more about ext4 and journaling.

### 3.20.3 References

| | |
|---|---|
| The kernel source | file:/usr/src/linux/fs/ext2/ |
| e2fsprogs (e2fsck) | http://e2fsprogs.sourceforge.net/ |
| Design & Implementation | http://e2fsprogs.sourceforge.net/ext2intro.html |
| Journaling (ext3) | ftp://ftp.uk.linux.org/pub/linux/sct/fs/jfs/ |
| Filesystem Resizing | http://ext2resize.sourceforge.net/ |
| Compression[1] | http://e2compr.sourceforge.net/ |

[1] no longer actively developed/supported (as of Apr 2001)

Implementations for:

| | |
|---|---|
| Windows 95/98/NT/2000 | http://www.chrysocome.net/explore2fs |
| Windows 95[Page 499, 1] | http://www.yipton.net/content.html#FSDEXT2 |
| DOS client[Page 499, 1] | ftp://metalab.unc.edu/pub/Linux/system/filesystems/ext2/ |
| OS/2[2] | ftp://metalab.unc.edu/pub/Linux/system/filesystems/ext2/ |
| RISC OS client | http://www.esw-heim.tu-clausthal.de/~marco/smorbrod/IscaFS/ |

## 3.21 Ext3 Filesystem

Ext3 was originally released in September 1999. Written by Stephen Tweedie for the 2.2 branch, and ported to 2.4 kernels by Peter Braam, Andreas Dilger, Andrew Morton, Alexander Viro, Ted Ts'o and Stephen Tweedie.

Ext3 is the ext2 filesystem enhanced with journalling capabilities. The filesystem is a subset of ext4 filesystem so use ext4 driver for accessing ext3 filesystems.

## 3.22 ext4 Data Structures and Algorithms

### 3.22.1 About this Book

This document attempts to describe the on-disk format for ext4 filesystems. The same general ideas should apply to ext2/3 filesystems as well, though they do not support all the features that ext4 supports, and the fields will be shorter.

**NOTE**: This is a work in progress, based on notes that the author (djwong) made while picking apart a filesystem by hand. The data structure definitions should be current as of Linux 4.18 and e2fsprogs-1.44. All comments and corrections are welcome, since there is undoubtedly plenty of lore that might not be reflected in freshly created demonstration filesystems.

#### License

This book is licensed under the terms of the GNU Public License, v2.

#### Terminology

ext4 divides a storage device into an array of logical blocks both to reduce bookkeeping overhead and to increase throughput by forcing larger transfer sizes. Generally, the block size will be 4KiB (the same size as pages on x86 and the block layer's default block size), though the actual size is calculated as 2 ^ (10 + `sb.s_log_block_size`) bytes. Throughout this document, disk locations are given in terms of these logical blocks, not raw LBAs, and not 1024-byte blocks. For the sake of convenience, the logical block size will be referred to as `$block_size` throughout the rest of the document.

[2] no longer actively developed/supported (as of Mar 2009)

When referenced in `preformatted text` blocks, `sb` refers to fields in the super block, and `inode` refers to fields in an inode table entry.

#### Other References

Also see https://www.nongnu.org/ext2-doc/ for quite a collection of information about ext2/3. Here's another old reference: http://wiki.osdev.org/Ext2

### 3.22.2 High Level Design

An ext4 file system is split into a series of block groups. To reduce performance difficulties due to fragmentation, the block allocator tries very hard to keep each file's blocks within the same group, thereby reducing seek times. The size of a block group is specified in `sb.s_blocks_per_group` blocks, though it can also calculated as 8 * `block_size_in_bytes`. With the default block size of 4KiB, each group will contain 32,768 blocks, for a length of 128MiB. The number of block groups is the size of the device divided by the size of a block group.

All fields in ext4 are written to disk in little-endian order. HOWEVER, all fields in jbd2 (the journal) are written to disk in big-endian order.

#### Blocks

ext4 allocates storage space in units of “blocks”. A block is a group of sectors between 1KiB and 64KiB, and the number of sectors must be an integral power of 2. Blocks are in turn grouped into larger units called block groups. Block size is specified at mkfs time and typically is 4KiB. You may experience mounting problems if block size is greater than page size (i.e. 64KiB blocks on a i386 which only has 4KiB memory pages). By default a filesystem can contain 2^32 blocks; if the '64bit' feature is enabled, then a filesystem can have 2^64 blocks. The location of structures is stored in terms of the block number the structure lives in and not the absolute offset on disk.

For 32-bit filesystems, limits are as follows:

| Item | 1KiB | 2KiB | 4KiB | 64KiB |
|---|---|---|---|---|
| Blocks | 2^32 | 2^32 | 2^32 | 2^32 |
| Inodes | 2^32 | 2^32 | 2^32 | 2^32 |
| File System Size | 4TiB | 8TiB | 16TiB | 256TiB |
| Blocks Per Block Group | 8,192 | 16,384 | 32,768 | 524,288 |
| Inodes Per Block Group | 8,192 | 16,384 | 32,768 | 524,288 |
| Block Group Size | 8MiB | 32MiB | 128MiB | 32GiB |
| Blocks Per File, Extents | 2^32 | 2^32 | 2^32 | 2^32 |
| Blocks Per File, Block Maps | 16,843,020 | 134,480,396 | 1,074,791,436 | 4,398,314,962,956 (really 2^32 due to field size limitations) |
| File Size, Extents | 4TiB | 8TiB | 16TiB | 256TiB |
| File Size, Block Maps | 16GiB | 256GiB | 4TiB | 256TiB |

For 64-bit filesystems, limits are as follows:

| Item | 1KiB | 2KiB | 4KiB | 64KiB |
|---|---|---|---|---|
| Blocks | 2^64 | 2^64 | 2^64 | 2^64 |
| Inodes | 2^32 | 2^32 | 2^32 | 2^32 |
| File System Size | 16ZiB | 32ZiB | 64ZiB | 1YiB |
| Blocks Per Block Group | 8,192 | 16,384 | 32,768 | 524,288 |
| Inodes Per Block Group | 8,192 | 16,384 | 32,768 | 524,288 |
| Block Group Size | 8MiB | 32MiB | 128MiB | 32GiB |
| Blocks Per File, Extents | 2^32 | 2^32 | 2^32 | 2^32 |
| Blocks Per File, Block Maps | 16,843,020 | 134,480,396 | 1,074,791,436 | 4,398,314,962,956 (really 2^32 due to field size limitations) |
| File Size, Extents | 4TiB | 8TiB | 16TiB | 256TiB |
| File Size, Block Maps | 16GiB | 256GiB | 4TiB | 256TiB |

Note: Files not using extents (i.e. files using block maps) must be placed within the first 2^32 blocks of a filesystem. Files with extents must be placed within the first 2^48 blocks of a filesystem. It's not clear what happens with larger filesystems.

### Layout

The layout of a standard block group is approximately as follows (each of these fields is discussed in a separate section below):

| Group 0 Padding | ext4 Super Block | Group Descriptors | Reserved GDT Blocks | Data Block Bitmap | inode Bitmap | inode Table | Data Blocks |
|---|---|---|---|---|---|---|---|
| 1024 bytes | 1 block | many blocks | many blocks | 1 block | 1 block | many blocks | many more blocks |

For the special case of block group 0, the first 1024 bytes are unused, to allow for the installation of x86 boot sectors and other oddities. The superblock will start at offset 1024 bytes, whichever block that happens to be (usually 0). However, if for some reason the block size = 1024, then block 0 is marked in use and the superblock goes in block 1. For all other block groups, there is no padding.

The ext4 driver primarily works with the superblock and the group descriptors that are found in block group 0. Redundant copies of the superblock and group descriptors are written to some of the block groups across the disk in case the beginning of the disk gets trashed, though not all block groups necessarily host a redundant copy (see following paragraph for more details). If the group does not have a redundant copy, the block group begins with the data block bitmap. Note also that when the filesystem is freshly formatted, mkfs will allocate "reserve GDT block" space after the block group descriptors and before the start of the block bitmaps to allow for future expansion of the filesystem. By default, a filesystem is allowed to increase in size by a factor of 1024x over the original filesystem size.

The location of the inode table is given by `grp.bg_inode_table_*`. It is continuous range of blocks large enough to contain `sb.s_inodes_per_group * sb.s_inode_size` bytes.

As for the ordering of items in a block group, it is generally established that the super block and the group descriptor table, if present, will be at the beginning of the block group. The bitmaps and the inode table can be anywhere, and it is quite possible for the bitmaps to come after the inode table, or for both to be in different groups (flex_bg). Leftover space is used for file data blocks, indirect block maps, extent tree blocks, and extended attributes.

### Flexible Block Groups

Starting in ext4, there is a new feature called flexible block groups (flex_bg). In a flex_bg, several block groups are tied together as one logical block group; the bitmap spaces and the inode table space in the first block group of the flex_bg are expanded to include the bitmaps and inode tables of all other block groups in the flex_bg. For example, if the flex_bg size is 4, then group 0 will contain (in order) the superblock, group descriptors, data block bitmaps for groups 0-3, inode bitmaps for groups 0-3, inode tables for groups 0-3, and the remaining space in group 0 is for file data. The effect of this is to group the block group metadata close together for faster loading, and to enable large files to be continuous on disk. Backup copies of the superblock and group descriptors are always at the beginning of block groups, even if flex_bg is enabled. The number of block groups that make up a flex_bg is given by `2 ^ sb.s_log_groups_per_flex`.

### Meta Block Groups

Without the option META_BG, for safety concerns, all block group descriptors copies are kept in the first block group. Given the default 128MiB(2^27 bytes) block group size and 64-byte group descriptors, ext4 can have at most 2^27/64 = 2^21 block groups. This limits the entire filesystem size to 2^21 * 2^27 = 2^48bytes or 256TiB.

The solution to this problem is to use the metablock group feature (META_BG), which is already in ext3 for all 2.6 releases. With the META_BG feature, ext4 filesystems are partitioned into many metablock groups. Each metablock group is a cluster of block groups whose group descriptor structures can be stored in a single disk block. For ext4 filesystems with 4 KB block size, a single metablock group partition includes 64 block groups, or 8 GiB of disk space. The metablock group feature moves the location of the group descriptors from the congested first block group of the whole filesystem into the first group of each metablock group itself. The backups are in the second and last group of each metablock group. This increases the 2^21 maximum block groups limit to the hard limit 2^32, allowing support for a 512PiB filesystem.

The change in the filesystem format replaces the current scheme where the superblock is followed by a variable-length set of block group descriptors. Instead, the superblock and a single block group descriptor block is placed at the beginning of the first, second, and last block groups in a meta-block group. A meta-block group is a collection of block groups which can be described by a single block group descriptor block. Since the size of the block group descriptor structure is 64 bytes, a meta-block group contains 16 block groups for filesystems with a 1KB block size, and 64 block groups for filesystems with a 4KB blocksize. Filesystems can either be created using this new block group descriptor layout, or existing filesystems can be resized on-line, and the field s_first_meta_bg in the superblock will indicate the first block group using this new layout.

Please see an important note about `BLOCK_UNINIT` in the section about block and inode bitmaps.

### Lazy Block Group Initialization

A new feature for ext4 are three block group descriptor flags that enable mkfs to skip initializing other parts of the block group metadata. Specifically, the INODE_UNINIT and BLOCK_UNINIT flags mean that the inode and block bitmaps for that group can be calculated and therefore the on-disk bitmap blocks are not initialized. This is generally the case for an empty block group or a block group containing only fixed-location block group metadata. The INODE_ZEROED flag means that the inode table has been initialized; mkfs will unset this flag and rely on the kernel to initialize the inode tables in the background.

By not writing zeroes to the bitmaps and inode table, mkfs time is reduced considerably. Note the feature flag is RO_COMPAT_GDT_CSUM, but the dumpe2fs output prints this as “uninit_bg”. They are the same thing.

### Special inodes

ext4 reserves some inode for special features, as follows:

| inode Number | Purpose |
|---|---|
| 0 | Doesn't exist; there is no inode 0. |
| 1 | List of defective blocks. |
| 2 | Root directory. |
| 3 | User quota. |
| 4 | Group quota. |
| 5 | Boot loader. |
| 6 | Undelete directory. |
| 7 | Reserved group descriptors inode. (“resize inode”) |
| 8 | Journal inode. |
| 9 | The “exclude” inode, for snapshots(?) |
| 10 | Replica inode, used for some non-upstream feature? |
| 11 | Traditional first non-reserved inode. Usually this is the lost+found directory. See s_first_ino in the superblock. |

Note that there are also some inodes allocated from non-reserved inode numbers for other filesystem features which are not referenced from standard directory hierarchy. These are generally reference from the superblock. They are:

| Superblock field | Description |
|---|---|
| s_lpf_ino | Inode number of lost+found directory. |
| s_prj_quota_inum | Inode number of quota file tracking project quotas |
| s_orphan_file_inum | Inode number of file tracking orphan inodes. |

### Block and Inode Allocation Policy

ext4 recognizes (better than ext3, anyway) that data locality is generally a desirably quality of a filesystem. On a spinning disk, keeping related blocks near each other reduces the amount of movement that the head actuator and disk must perform to access a data block, thus speeding up disk IO. On an SSD there of course are no moving parts, but locality can increase the size of each transfer request while reducing the total number of requests. This locality may also have the effect of concentrating writes on a single erase block, which can speed up file rewrites significantly. Therefore, it is useful to reduce fragmentation whenever possible.

The first tool that ext4 uses to combat fragmentation is the multi-block allocator. When a file is first created, the block allocator speculatively allocates 8KiB of disk space to the file on the assumption that the space will get written soon. When the file is closed, the unused speculative allocations are of course freed, but if the speculation is correct (typically the case for full writes of small files) then the file data gets written out in a single multi-block extent. A second related trick that ext4 uses is delayed allocation. Under this scheme, when a file needs more blocks to absorb file writes, the filesystem defers deciding the exact placement on the disk until all the dirty buffers are being written out to disk. By not committing to a particular placement until

it's absolutely necessary (the commit timeout is hit, or sync() is called, or the kernel runs out of memory), the hope is that the filesystem can make better location decisions.

The third trick that ext4 (and ext3) uses is that it tries to keep a file's data blocks in the same block group as its inode. This cuts down on the seek penalty when the filesystem first has to read a file's inode to learn where the file's data blocks live and then seek over to the file's data blocks to begin I/O operations.

The fourth trick is that all the inodes in a directory are placed in the same block group as the directory, when feasible. The working assumption here is that all the files in a directory might be related, therefore it is useful to try to keep them all together.

The fifth trick is that the disk volume is cut up into 128MB block groups; these mini-containers are used as outlined above to try to maintain data locality. However, there is a deliberate quirk -- when a directory is created in the root directory, the inode allocator scans the block groups and puts that directory into the least heavily loaded block group that it can find. This encourages directories to spread out over a disk; as the top-level directory/file blobs fill up one block group, the allocators simply move on to the next block group. Allegedly this scheme evens out the loading on the block groups, though the author suspects that the directories which are so unlucky as to land towards the end of a spinning drive get a raw deal performance-wise.

Of course if all of these mechanisms fail, one can always use e4defrag to defragment files.

### Checksums

Starting in early 2012, metadata checksums were added to all major ext4 and jbd2 data structures. The associated feature flag is metadata_csum. The desired checksum algorithm is indicated in the superblock, though as of October 2012 the only supported algorithm is crc32c. Some data structures did not have space to fit a full 32-bit checksum, so only the lower 16 bits are stored. Enabling the 64bit feature increases the data structure size so that full 32-bit checksums can be stored for many data structures. However, existing 32-bit filesystems cannot be extended to enable 64bit mode, at least not without the experimental resize2fs patches to do so.

Existing filesystems can have checksumming added by running `tune2fs -O metadata_csum` against the underlying device. If tune2fs encounters directory blocks that lack sufficient empty space to add a checksum, it will request that you run `e2fsck -D` to have the directories rebuilt with checksums. This has the added benefit of removing slack space from the directory files and rebalancing the htree indexes. If you _ignore_ this step, your directories will not be protected by a checksum!

The following table describes the data elements that go into each type of checksum. The checksum function is whatever the superblock describes (crc32c as of October 2013) unless noted otherwise.

| Metadata | Length | Ingredients |
|---|---|---|
| Superblock | __le32 | The entire superblock up to the checksum field. The UUID lives inside the superblock. |
| MMP | __le32 | UUID + the entire MMP block up to the checksum field. |
| Extended Attributes | __le32 | UUID + the entire extended attribute block. The checksum field is set to zero. |
| Directory Entries | __le32 | UUID + inode number + inode generation + the directory block up to the fake entry enclosing the checksum field. |
| HTREE Nodes | __le32 | UUID + inode number + inode generation + all valid extents + HTREE tail. The checksum field is set to zero. |
| Extents | __le32 | UUID + inode number + inode generation + the entire extent block up to the checksum field. |
| Bitmaps | __le32 or __le16 | UUID + the entire bitmap. Checksums are stored in the group descriptor, and truncated if the group descriptor size is 32 bytes (i.e. ^64bit) |
| Inodes | __le32 | UUID + inode number + inode generation + the entire inode. The checksum field is set to zero. Each inode has its own checksum. |
| Group Descriptors | __le16 | If metadata_csum, then UUID + group number + the entire descriptor; else if gdt_csum, then crc16(UUID + group number + the entire descriptor). In all cases, only the lower 16 bits are stored. |

### Bigalloc

At the moment, the default size of a block is 4KiB, which is a commonly supported page size on most MMU-capable hardware. This is fortunate, as ext4 code is not prepared to handle the case where the block size exceeds the page size. However, for a filesystem of mostly huge files, it is desirable to be able to allocate disk blocks in units of multiple blocks to reduce both fragmentation and metadata overhead. The bigalloc feature provides exactly this ability.

The bigalloc feature (EXT4_FEATURE_RO_COMPAT_BIGALLOC) changes ext4 to use clustered allocation, so that each bit in the ext4 block allocation bitmap addresses a power of two number of blocks. For example, if the file system is mainly going to be storing large files in the 4-32 megabyte range, it might make sense to set a cluster size of 1 megabyte. This means that each bit in the block allocation bitmap now addresses 256 4k blocks. This shrinks the total size of the block allocation bitmaps for a 2T file system from 64 megabytes to 256 kilobytes. It also means that a block group addresses 32 gigabytes instead of 128 megabytes, also shrinking the amount of file system overhead for metadata.

The administrator can set a block cluster size at mkfs time (which is stored in the s_log_cluster_size field in the superblock); from then on, the block bitmaps track clusters, not individual blocks. This means that block groups can be several gigabytes in size (instead of just 128MiB); however, the minimum allocation unit becomes a cluster, not a block, even for directories. TaoBao had a patchset to extend the “use units of clusters instead of blocks” to the extent tree, though it is not clear where those patches went-- they eventually morphed into “extent tree v2” but that code has not landed as of May 2015.

### Inline Data

The inline data feature was designed to handle the case that a file's data is so tiny that it readily fits inside the inode, which (theoretically) reduces disk block consumption and reduces seeks. If the file is smaller than 60 bytes, then the data are stored inline in `inode.i_block`. If the rest of the file would fit inside the extended attribute space, then it might be found as an extended attribute "system.data" within the inode body ("ibody EA"). This of course constrains the amount of extended attributes one can attach to an inode. If the data size increases beyond i_block + ibody EA, a regular block is allocated and the contents moved to that block.

Pending a change to compact the extended attribute key used to store inline data, one ought to be able to store 160 bytes of data in a 256-byte inode (as of June 2015, when i_extra_isize is 28). Prior to that, the limit was 156 bytes due to inefficient use of inode space.

The inline data feature requires the presence of an extended attribute for "system.data", even if the attribute value is zero length.

### Inline Directories

The first four bytes of i_block are the inode number of the parent directory. Following that is a 56-byte space for an array of directory entries; see `struct ext4_dir_entry`. If there is a "system.data" attribute in the inode body, the EA value is an array of `struct ext4_dir_entry` as well. Note that for inline directories, the i_block and EA space are treated as separate dirent blocks; directory entries cannot span the two.

Inline directory entries are not checksummed, as the inode checksum should protect all inline data contents.

### Large Extended Attribute Values

To enable ext4 to store extended attribute values that do not fit in the inode or in the single extended attribute block attached to an inode, the EA_INODE feature allows us to store the value in the data blocks of a regular file inode. This "EA inode" is linked only from the extended attribute name index and must not appear in a directory entry. The inode's i_atime field is used to store a checksum of the xattr value; and i_ctime/i_version store a 64-bit reference count, which enables sharing of large xattr values between multiple owning inodes. For backward compatibility with older versions of this feature, the i_mtime/i_generation *may* store a back-reference to the inode number and i_generation of the **one** owning inode (in cases where the EA inode is not referenced by multiple inodes) to verify that the EA inode is the correct one being accessed.

### Verity files

ext4 supports fs-verity, which is a filesystem feature that provides Merkle tree based hashing for individual readonly files. Most of fs-verity is common to all filesystems that support it; see *Documentation/filesystems/fsverity.rst* for the fs-verity documentation. However, the on-disk layout of the verity metadata is filesystem-specific. On ext4, the verity metadata is stored after the end of the file data itself, in the following format:

- Zero-padding to the next 65536-byte boundary. This padding need not actually be allocated on-disk, i.e. it may be a hole.

- The Merkle tree, as documented in *Documentation/filesystems/fsverity.rst*, with the tree levels stored in order from root to leaf, and the tree blocks within each level stored in their natural order.
- Zero-padding to the next filesystem block boundary.
- The verity descriptor, as documented in *Documentation/filesystems/fsverity.rst*, with optionally appended signature blob.
- Zero-padding to the next offset that is 4 bytes before a filesystem block boundary.
- The size of the verity descriptor in bytes, as a 4-byte little endian integer.

Verity inodes have EXT4_VERITY_FL set, and they must use extents, i.e. EXT4_EXTENTS_FL must be set and EXT4_INLINE_DATA_FL must be clear. They can have EXT4_ENCRYPT_FL set, in which case the verity metadata is encrypted as well as the data itself.

Verity files cannot have blocks allocated past the end of the verity metadata.

Verity and DAX are not compatible and attempts to set both of these flags on a file will fail.

### 3.22.3 Global Structures

The filesystem is sharded into a number of block groups, each of which have static metadata at fixed locations.

#### Super Block

The superblock records various information about the enclosing filesystem, such as block counts, inode counts, supported features, maintenance information, and more.

If the sparse_super feature flag is set, redundant copies of the superblock and group descriptors are kept only in the groups whose group number is either 0 or a power of 3, 5, or 7. If the flag is not set, redundant copies are kept in all groups.

The superblock checksum is calculated against the superblock structure, which includes the FS UUID.

The ext4 superblock is laid out as follows in `struct ext4_super_block`:

| Offset | Size | Name | Description |
|---|---|---|---|
| 0x0 | __le32 | s_inodes_count | Total inode count. |
| 0x4 | __le32 | s_blocks_count_lo | Total block count. |
| 0x8 | __le32 | s_r_blocks_count_lo | This number of blocks can only be allocated by the super-user. |
| 0xC | __le32 | s_free_blocks_count_lo | Free block count. |
| 0x10 | __le32 | s_free_inodes_count | Free inode count. |
| 0x14 | __le32 | s_first_data_block | First data block. This must be at least 1 for 1k-block filesystems and is typically 0 for all other block sizes. |
| 0x18 | __le32 | s_log_block_size | Block size is 2 ^ (10 + s_log_block_size). |

continues on next page

Table 1 – continued from previous page

| Offset | Size | Name | Description |
|---|---|---|---|
| 0x1C | __le32 | s_log_cluster_size | Cluster size is 2 ^ (10 + s_log_cluster_size) blocks if bigalloc is enabled. Otherwise s_log_cluster_size must equal s_log_block_size. |
| 0x20 | __le32 | s_blocks_per_group | Blocks per group. |
| 0x24 | __le32 | s_clusters_per_group | Clusters per group, if bigalloc is enabled. Otherwise s_clusters_per_group must equal s_blocks_per_group. |
| 0x28 | __le32 | s_inodes_per_group | Inodes per group. |
| 0x2C | __le32 | s_mtime | Mount time, in seconds since the epoch. |
| 0x30 | __le32 | s_wtime | Write time, in seconds since the epoch. |
| 0x34 | __le16 | s_mnt_count | Number of mounts since the last fsck. |
| 0x36 | __le16 | s_max_mnt_count | Number of mounts beyond which a fsck is needed. |
| 0x38 | __le16 | s_magic | Magic signature, 0xEF53 |
| 0x3A | __le16 | s_state | File system state. See *super_state* for more info. |
| 0x3C | __le16 | s_errors | Behaviour when detecting errors. See *super_errors* for more info. |
| 0x3E | __le16 | s_minor_rev_level | Minor revision level. |
| 0x40 | __le32 | s_lastcheck | Time of last check, in seconds since the epoch. |
| 0x44 | __le32 | s_checkinterval | Maximum time between checks, in seconds. |
| 0x48 | __le32 | s_creator_os | Creator OS. See the table *super_creator* for more info. |
| 0x4C | __le32 | s_rev_level | Revision level. See the table *super_revision* for more info. |
| 0x50 | __le16 | s_def_resuid | Default uid for reserved blocks. |
| 0x52 | __le16 | s_def_resgid | Default gid for reserved blocks. |
| | | | These fields are for EXT4_DYNAMIC_REV superblocks only.<br>Note: the difference between the compatible feature set and the incompatible feature set is that if there is a bit set in the incompatible feature set that the kernel doesn't know about, it should refuse to mount the filesystem.<br>e2fsck's requirements are more strict; if it doesn't know about a feature in either the compatible or incompatible feature set, it must abort and not try to meddle with things it doesn't understand... |
| 0x54 | __le32 | s_first_ino | First non-reserved inode. |
| 0x58 | __le16 | s_inode_size | Size of inode structure, in bytes. |
| 0x5A | __le16 | s_block_group_nr | Block group # of this superblock. |

continues on next page

Table 1 – continued from previous page

| Offset | Size | Name | Description |
|---|---|---|---|
| 0x5C | __le32 | s_feature_compat | Compatible feature set flags. Kernel can still read/write this fs even if it doesn't understand a flag; fsck should not do that. See the *super_compat* table for more info. |
| 0x60 | __le32 | s_feature_incompat | Incompatible feature set. If the kernel or fsck doesn't understand one of these bits, it should stop. See the *super_incompat* table for more info. |
| 0x64 | __le32 | s_feature_ro_compat | Readonly-compatible feature set. If the kernel doesn't understand one of these bits, it can still mount read-only. See the *super_rocompat* table for more info. |
| 0x68 | __u8 | s_uuid[16] | 128-bit UUID for volume. |
| 0x78 | char | s_volume_name[16] | Volume label. |
| 0x88 | char | s_last_mounted[64] | Directory where filesystem was last mounted. |
| 0xC8 | __le32 | s_algorithm_usage_bitmap | For compression (Not used in e2fsprogs/Linux) |
| | | | Performance hints. Directory preallocation should only happen if the EXT4_FEATURE_COMPAT_DIR_PREALLOC flag is on. |
| 0xCC | __u8 | s_prealloc_blocks | #. of blocks to try to preallocate for ... files? (Not used in e2fsprogs/Linux) |
| 0xCD | __u8 | s_prealloc_dir_blocks | #. of blocks to preallocate for directories. (Not used in e2fsprogs/Linux) |
| 0xCE | __le16 | s_reserved_gdt_blocks | Number of reserved GDT entries for future filesystem expansion. |
| | | | Journalling support is valid only if EXT4_FEATURE_COMPAT_HAS_JOURNAL is set. |
| 0xD0 | __u8 | s_journal_uuid[16] | UUID of journal superblock |
| 0xE0 | __le32 | s_journal_inum | inode number of journal file. |
| 0xE4 | __le32 | s_journal_dev | Device number of journal file, if the external journal feature flag is set. |
| 0xE8 | __le32 | s_last_orphan | Start of list of orphaned inodes to delete. |
| 0xEC | __le32 | s_hash_seed[4] | HTREE hash seed. |
| 0xFC | __u8 | s_def_hash_version | Default hash algorithm to use for directory hashes. See *super_def_hash* for more info. |
| 0xFD | __u8 | s_jnl_backup_type | If this value is 0 or EXT3_JNL_BACKUP_BLOCKS (1), then the `s_jnl_blocks` field contains a duplicate copy of the inode's `i_block[]` array and `i_size`. |
| 0xFE | __le16 | s_desc_size | Size of group descriptors, in bytes, if the 64bit incompat feature flag is set. |
| 0x100 | __le32 | s_default_mount_opts | Default mount options. See the *super_mountopts* table for more info. |

continues on next page

Table 1 – continued from previous page

| Offset | Size | Name | Description |
|---|---|---|---|
| 0x104 | __le32 | s_first_meta_bg | First metablock block group, if the meta_bg feature is enabled. |
| 0x108 | __le32 | s_mkfs_time | When the filesystem was created, in seconds since the epoch. |
| 0x10C | __le32 | s_jnl_blocks[17] | Backup copy of the journal inode's `i_block[]` array in the first 15 elements and i_size_high and i_size in the 16th and 17th elements, respectively. |
| | | | 64bit support is valid only if EXT4_FEATURE_COMPAT_64BIT is set. |
| 0x150 | __le32 | s_blocks_count_hi | High 32-bits of the block count. |
| 0x154 | __le32 | s_r_blocks_count_hi | High 32-bits of the reserved block count. |
| 0x158 | __le32 | s_free_blocks_count_hi | High 32-bits of the free block count. |
| 0x15C | __le16 | s_min_extra_isize | All inodes have at least # bytes. |
| 0x15E | __le16 | s_want_extra_isize | New inodes should reserve # bytes. |
| 0x160 | __le32 | s_flags | Miscellaneous flags. See the *super_flags* table for more info. |
| 0x164 | __le16 | s_raid_stride | RAID stride. This is the number of logical blocks read from or written to the disk before moving to the next disk. This affects the placement of filesystem metadata, which will hopefully make RAID storage faster. |
| 0x166 | __le16 | s_mmp_interval | #. seconds to wait in multi-mount prevention (MMP) checking. In theory, MMP is a mechanism to record in the superblock which host and device have mounted the filesystem, in order to prevent multiple mounts. This feature does not seem to be implemented... |
| 0x168 | __le64 | s_mmp_block | Block # for multi-mount protection data. |
| 0x170 | __le32 | s_raid_stripe_width | RAID stripe width. This is the number of logical blocks read from or written to the disk before coming back to the current disk. This is used by the block allocator to try to reduce the number of read-modify-write operations in a RAID5/6. |
| 0x174 | __u8 | s_log_groups_per_flex | Size of a flexible block group is 2 ^ `s_log_groups_per_flex`. |
| 0x175 | __u8 | s_checksum_type | Metadata checksum algorithm type. The only valid value is 1 (crc32c). |
| 0x176 | __le16 | s_reserved_pad | |
| 0x178 | __le64 | s_kbytes_written | Number of KiB written to this filesystem over its lifetime. |
| 0x180 | __le32 | s_snapshot_inum | inode number of active snapshot. (Not used in e2fsprogs/Linux.) |
| 0x184 | __le32 | s_snapshot_id | Sequential ID of active snapshot. (Not used in e2fsprogs/Linux.) |

continues on next page

Table 1 – continued from previous page

| Offset | Size | Name | Description |
|---|---|---|---|
| 0x188 | __le64 | s_snapshot_r_blocks_count | Number of blocks reserved for active snapshot's future use. (Not used in e2fsprogs/Linux.) |
| 0x190 | __le32 | s_snapshot_list | inode number of the head of the on-disk snapshot list. (Not used in e2fsprogs/Linux.) |
| 0x194 | __le32 | s_error_count | Number of errors seen. |
| 0x198 | __le32 | s_first_error_time | First time an error happened, in seconds since the epoch. |
| 0x19C | __le32 | s_first_error_ino | inode involved in first error. |
| 0x1A0 | __le64 | s_first_error_block | Number of block involved of first error. |
| 0x1A8 | __u8 | s_first_error_func[32] | Name of function where the error happened. |
| 0x1C8 | __le32 | s_first_error_line | Line number where error happened. |
| 0x1CC | __le32 | s_last_error_time | Time of most recent error, in seconds since the epoch. |
| 0x1D0 | __le32 | s_last_error_ino | inode involved in most recent error. |
| 0x1D4 | __le32 | s_last_error_line | Line number where most recent error happened. |
| 0x1D8 | __le64 | s_last_error_block | Number of block involved in most recent error. |
| 0x1E0 | __u8 | s_last_error_func[32] | Name of function where the most recent error happened. |
| 0x200 | __u8 | s_mount_opts[64] | ASCIIZ string of mount options. |
| 0x240 | __le32 | s_usr_quota_inum | Inode number of user quota file. |
| 0x244 | __le32 | s_grp_quota_inum | Inode number of group quota file. |
| 0x248 | __le32 | s_overhead_blocks | Overhead blocks/clusters in fs. (Huh? This field is always zero, which means that the kernel calculates it dynamically.) |
| 0x24C | __le32 | s_backup_bgs[2] | Block groups containing superblock backups (if sparse_super2) |
| 0x254 | __u8 | s_encrypt_algos[4] | Encryption algorithms in use. There can be up to four algorithms in use at any time; valid algorithm codes are given in the *super_encrypt* table below. |
| 0x258 | __u8 | s_encrypt_pw_salt[16] | Salt for the string2key algorithm for encryption. |
| 0x268 | __le32 | s_lpf_ino | Inode number of lost+found |
| 0x26C | __le32 | s_prj_quota_inum | Inode that tracks project quotas. |
| 0x270 | __le32 | s_checksum_seed | Checksum seed used for metadata_csum calculations. This value is crc32c(~0, $orig_fs_uuid). |
| 0x274 | __u8 | s_wtime_hi | Upper 8 bits of the s_wtime field. |
| 0x275 | __u8 | s_mtime_hi | Upper 8 bits of the s_mtime field. |
| 0x276 | __u8 | s_mkfs_time_hi | Upper 8 bits of the s_mkfs_time field. |
| 0x277 | __u8 | s_lastcheck_hi | Upper 8 bits of the s_lastcheck field. |
| 0x278 | __u8 | s_first_error_time_hi | Upper 8 bits of the s_first_error_time field. |
| 0x279 | __u8 | s_last_error_time_hi | Upper 8 bits of the s_last_error_time field. |
| 0x27A | __u8 | s_pad[2] | Zero padding. |

continues on next page

Table 1 – continued from previous page

| Offset | Size | Name | Description |
|---|---|---|---|
| 0x27C | __le16 | s_encoding | Filename charset encoding. |
| 0x27E | __le16 | s_encoding_flags | Filename charset encoding flags. |
| 0x280 | __le32 | s_orphan_file_inum | Orphan file inode number. |
| 0x284 | __le32 | s_reserved[94] | Padding to the end of the block. |
| 0x3FC | __le32 | s_checksum | Superblock checksum. |

The superblock state is some combination of the following:

| Value | Description |
|---|---|
| 0x0001 | Cleanly umounted |
| 0x0002 | Errors detected |
| 0x0004 | Orphans being recovered |

The superblock error policy is one of the following:

| Value | Description |
|---|---|
| 1 | Continue |
| 2 | Remount read-only |
| 3 | Panic |

The filesystem creator is one of the following:

| Value | Description |
|---|---|
| 0 | Linux |
| 1 | Hurd |
| 2 | Masix |
| 3 | FreeBSD |
| 4 | Lites |

The superblock revision is one of the following:

| Value | Description |
|---|---|
| 0 | Original format |
| 1 | v2 format w/ dynamic inode sizes |

Note that `EXT4_DYNAMIC_REV` refers to a revision 1 or newer filesystem.

The superblock compatible features field is a combination of any of the following:

| Value | Description |
|---|---|
| 0x1 | Directory preallocation (COMPAT_DIR_PREALLOC). |
| 0x2 | “imagic inodes”. Not clear from the code what this does (COMPAT_IMAGIC_INODES). |
| 0x4 | Has a journal (COMPAT_HAS_JOURNAL). |
| 0x8 | Supports extended attributes (COMPAT_EXT_ATTR). |
| 0x10 | Has reserved GDT blocks for filesystem expansion (COMPAT_RESIZE_INODE). Requires RO_COMPAT_SPARSE_SUPER. |
| 0x20 | Has directory indices (COMPAT_DIR_INDEX). |
| 0x40 | “Lazy BG”. Not in Linux kernel, seems to have been for uninitialized block groups? (COMPAT_LAZY_BG) |
| 0x80 | “Exclude inode”. Not used. (COMPAT_EXCLUDE_INODE). |
| 0x100 | “Exclude bitmap”. Seems to be used to indicate the presence of snapshot-related exclude bitmaps? Not defined in kernel or used in e2fsprogs (COMPAT_EXCLUDE_BITMAP). |
| 0x200 | Sparse Super Block, v2. If this flag is set, the SB field s_backup_bgs points to the two block groups that contain backup superblocks (COMPAT_SPARSE_SUPER2). |
| 0x400 | Fast commits supported. Although fast commits blocks are backward incompatible, fast commit blocks are not always present in the journal. If fast commit blocks are present in the journal, JBD2 incompat feature (JBD2_FEATURE_INCOMPAT_FAST_COMMIT) gets set (COMPAT_FAST_COMMIT). |
| 0x1000 | Orphan file allocated. This is the special file for more efficient tracking of unlinked but still open inodes. When there may be any entries in the file, we additionally set proper rocompat feature (RO_COMPAT_ORPHAN_PRESENT). |

The superblock incompatible features field is a combination of any of the following:

| Value | Description |
|---|---|
| 0x1 | Compression (INCOMPAT_COMPRESSION). |
| 0x2 | Directory entries record the file type. See ext4_dir_entry_2 below (INCOMPAT_FILETYPE). |
| 0x4 | Filesystem needs recovery (INCOMPAT_RECOVER). |
| 0x8 | Filesystem has a separate journal device (INCOMPAT_JOURNAL_DEV). |
| 0x10 | Meta block groups. See the earlier discussion of this feature (INCOMPAT_META_BG). |
| 0x40 | Files in this filesystem use extents (INCOMPAT_EXTENTS). |
| 0x80 | Enable a filesystem size of 2^64 blocks (INCOMPAT_64BIT). |
| 0x100 | Multiple mount protection (INCOMPAT_MMP). |
| 0x200 | Flexible block groups. See the earlier discussion of this feature (INCOMPAT_FLEX_BG). |
| 0x400 | Inodes can be used to store large extended attribute values (INCOMPAT_EA_INODE). |
| 0x1000 | Data in directory entry (INCOMPAT_DIRDATA). (Not implemented?) |
| 0x2000 | Metadata checksum seed is stored in the superblock. This feature enables the administrator to change the UUID of a metadata_csum filesystem while the filesystem is mounted; without it, the checksum definition requires all metadata blocks to be rewritten (INCOMPAT_CSUM_SEED). |
| 0x4000 | Large directory >2GB or 3-level htree (INCOMPAT_LARGEDIR). Prior to this feature, directories could not be larger than 4GiB and could not have an htree more than 2 levels deep. If this feature is enabled, directories can be larger than 4GiB and have a maximum htree depth of 3. |
| 0x8000 | Data in inode (INCOMPAT_INLINE_DATA). |
| 0x10000 | Encrypted inodes are present on the filesystem. (INCOMPAT_ENCRYPT). |

The superblock read-only compatible features field is a combination of any of the following:

| Value | Description |
|---|---|
| 0x1 | Sparse superblocks. See the earlier discussion of this feature (RO_COMPAT_SPARSE_SUPER). |
| 0x2 | This filesystem has been used to store a file greater than 2GiB (RO_COMPAT_LARGE_FILE). |
| 0x4 | Not used in kernel or e2fsprogs (RO_COMPAT_BTREE_DIR). |
| 0x8 | This filesystem has files whose sizes are represented in units of logical blocks, not 512-byte sectors. This implies a very large file indeed! (RO_COMPAT_HUGE_FILE) |
| 0x10 | Group descriptors have checksums. In addition to detecting corruption, this is useful for lazy formatting with uninitialized groups (RO_COMPAT_GDT_CSUM). |
| 0x20 | Indicates that the old ext3 32,000 subdirectory limit no longer applies (RO_COMPAT_DIR_NLINK). A directory's i_links_count will be set to 1 if it is incremented past 64,999. |
| 0x40 | Indicates that large inodes exist on this filesystem (RO_COMPAT_EXTRA_ISIZE). |
| 0x80 | This filesystem has a snapshot (RO_COMPAT_HAS_SNAPSHOT). |
| 0x100 | Quota (RO_COMPAT_QUOTA). |
| 0x200 | This filesystem supports "bigalloc", which means that file extents are tracked in units of clusters (of blocks) instead of blocks (RO_COMPAT_BIGALLOC). |
| 0x400 | This filesystem supports metadata checksumming. (RO_COMPAT_METADATA_CSUM; implies RO_COMPAT_GDT_CSUM, though GDT_CSUM must not be set) |
| 0x800 | Filesystem supports replicas. This feature is neither in the kernel nor e2fsprogs. (RO_COMPAT_REPLICA) |
| 0x1000 | Read-only filesystem image; the kernel will not mount this image read-write and most tools will refuse to write to the image. (RO_COMPAT_READONLY) |
| 0x2000 | Filesystem tracks project quotas. (RO_COMPAT_PROJECT) |
| 0x8000 | Verity inodes may be present on the filesystem. (RO_COMPAT_VERITY) |
| 0x10000 | Indicates orphan file may have valid orphan entries and thus we need to clean them up when mounting the filesystem (RO_COMPAT_ORPHAN_PRESENT). |

The `s_def_hash_version` field is one of the following:

| Value | Description |
|---|---|
| 0x0 | Legacy. |
| 0x1 | Half MD4. |
| 0x2 | Tea. |
| 0x3 | Legacy, unsigned. |
| 0x4 | Half MD4, unsigned. |
| 0x5 | Tea, unsigned. |

The `s_default_mount_opts` field is any combination of the following:

| Value | Description |
|---|---|
| 0x0001 | Print debugging info upon (re)mount. (EXT4_DEFM_DEBUG) |
| 0x0002 | New files take the gid of the containing directory (instead of the fsgid of the current process). (EXT4_DEFM_BSDGROUPS) |
| 0x0004 | Support userspace-provided extended attributes. (EXT4_DEFM_XATTR_USER) |
| 0x0008 | Support POSIX access control lists (ACLs). (EXT4_DEFM_ACL) |
| 0x0010 | Do not support 32-bit UIDs. (EXT4_DEFM_UID16) |
| 0x0020 | All data and metadata are committed to the journal. (EXT4_DEFM_JMODE_DATA) |
| 0x0040 | All data are flushed to the disk before metadata are committed to the journal. (EXT4_DEFM_JMODE_ORDERED) |
| 0x0060 | Data ordering is not preserved; data may be written after the metadata has been written. (EXT4_DEFM_JMODE_WBACK) |
| 0x0100 | Disable write flushes. (EXT4_DEFM_NOBARRIER) |
| 0x0200 | Track which blocks in a filesystem are metadata and therefore should not be used as data blocks. This option will be enabled by default on 3.18, hopefully. (EXT4_DEFM_BLOCK_VALIDITY) |
| 0x0400 | Enable DISCARD support, where the storage device is told about blocks becoming unused. (EXT4_DEFM_DISCARD) |
| 0x0800 | Disable delayed allocation. (EXT4_DEFM_NODELALLOC) |

The `s_flags` field is any combination of the following:

| Value | Description |
|---|---|
| 0x0001 | Signed directory hash in use. |
| 0x0002 | Unsigned directory hash in use. |
| 0x0004 | To test development code. |

The `s_encrypt_algos` list can contain any of the following:

| Value | Description |
|---|---|
| 0 | Invalid algorithm (ENCRYPTION_MODE_INVALID). |
| 1 | 256-bit AES in XTS mode (ENCRYPTION_MODE_AES_256_XTS). |
| 2 | 256-bit AES in GCM mode (ENCRYPTION_MODE_AES_256_GCM). |
| 3 | 256-bit AES in CBC mode (ENCRYPTION_MODE_AES_256_CBC). |

Total size of the superblock is 1024 bytes.

### Block Group Descriptors

Each block group on the filesystem has one of these descriptors associated with it. As noted in the Layout section above, the group descriptors (if present) are the second item in the block group. The standard configuration is for each block group to contain a full copy of the block group descriptor table unless the sparse_super feature flag is set.

Notice how the group descriptor records the location of both bitmaps and the inode table (i.e. they can float). This means that within a block group, the only data structures with fixed locations are the superblock and the group descriptor table. The flex_bg mechanism uses this property to group several block groups into a flex group and lay out all of the groups' bitmaps and inode tables into one long run in the first group of the flex group.

If the meta_bg feature flag is set, then several block groups are grouped together into a meta group. Note that in the meta_bg case, however, the first and last two block groups within the larger meta group contain only group descriptors for the groups inside the meta group.

flex_bg and meta_bg do not appear to be mutually exclusive features.

In ext2, ext3, and ext4 (when the 64bit feature is not enabled), the block group descriptor was only 32 bytes long and therefore ends at bg_checksum. On an ext4 filesystem with the 64bit feature enabled, the block group descriptor expands to at least the 64 bytes described below; the size is stored in the superblock.

If gdt_csum is set and metadata_csum is not set, the block group checksum is the crc16 of the FS UUID, the group number, and the group descriptor structure. If metadata_csum is set, then the block group checksum is the lower 16 bits of the checksum of the FS UUID, the group number, and the group descriptor structure. Both block and inode bitmap checksums are calculated against the FS UUID, the group number, and the entire bitmap.

The block group descriptor is laid out in `struct ext4_group_desc`.

| Offset | Size | Name | Description |
|---|---|---|---|
| 0x0 | __le32 | bg_block_bitmap_lo | Lower 32-bits of location of block bitmap. |
| 0x4 | __le32 | bg_inode_bitmap_lo | Lower 32-bits of location of inode bitmap. |
| 0x8 | __le32 | bg_inode_table_lo | Lower 32-bits of location of inode table. |
| 0xC | __le16 | bg_free_blocks_count_lo | Lower 16-bits of free block count. |
| 0xE | __le16 | bg_free_inodes_count_lo | Lower 16-bits of free inode count. |
| 0x10 | __le16 | bg_used_dirs_count_lo | Lower 16-bits of directory count. |
| 0x12 | __le16 | bg_flags | Block group flags. See the *bgflags* table below. |
| 0x14 | __le32 | bg_exclude_bitmap_lo | Lower 32-bits of location of snapshot exclusion bitmap. |
| 0x18 | __le16 | bg_block_bitmap_csum_lo | Lower 16-bits of the block bitmap checksum. |
| 0x1A | __le16 | bg_inode_bitmap_csum_lo | Lower 16-bits of the inode bitmap checksum. |
| 0x1C | __le16 | bg_itable_unused_lo | Lower 16-bits of unused inode count. If set, we needn't scan past the `(sb.s_inodes_per_group - gdt. bg_itable_unused)` th entry in the inode table for this group. |
| 0x1E | __le16 | bg_checksum | Group descriptor checksum; crc16(sb_uuid+group_num+bg_desc) if the RO_COMPAT_GDT_CSUM feature is set, or crc32c(sb_uuid+group_num+bg_desc) & 0xFFFF if the RO_COMPAT_METADATA_CSUM feature is set. The bg_checksum field in bg_desc is skipped when calculating crc16 checksum, and set to zero if crc32c checksum is used. |
| | | | These fields only exist if the 64bit feature is enabled and s_desc_size > 32. |
| 0x20 | __le32 | bg_block_bitmap_hi | Upper 32-bits of location of block bitmap. |
| 0x24 | __le32 | bg_inode_bitmap_hi | Upper 32-bits of location of inodes bitmap. |
| 0x28 | __le32 | bg_inode_table_hi | Upper 32-bits of location of inodes table. |
| 0x2C | __le16 | bg_free_blocks_count_hi | Upper 16-bits of free block count. |
| 0x2E | __le16 | bg_free_inodes_count_hi | Upper 16-bits of free inode count. |
| 0x30 | __le16 | bg_used_dirs_count_hi | Upper 16-bits of directory count. |
| 0x32 | __le16 | bg_itable_unused_hi | Upper 16-bits of unused inode count. |
| 0x34 | __le32 | bg_exclude_bitmap_hi | Upper 32-bits of location of snapshot exclusion bitmap. |
| 0x38 | __le16 | bg_block_bitmap_csum_hi | Upper 16-bits of the block bitmap checksum. |
| 0x3A | __le16 | bg_inode_bitmap_csum_hi | Upper 16-bits of the inode bitmap checksum. |
| 0x3C | __u32 | bg_reserved | Padding to 64 bytes. |

Block group flags can be any combination of the following:

| Value | Description |
|---|---|
| 0x1 | inode table and bitmap are not initialized (EXT4_BG_INODE_UNINIT). |
| 0x2 | block bitmap is not initialized (EXT4_BG_BLOCK_UNINIT). |
| 0x4 | inode table is zeroed (EXT4_BG_INODE_ZEROED). |

### Block and inode Bitmaps

The data block bitmap tracks the usage of data blocks within the block group.

The inode bitmap records which entries in the inode table are in use.

As with most bitmaps, one bit represents the usage status of one data block or inode table entry. This implies a block group size of 8 * number_of_bytes_in_a_logical_block.

NOTE: If `BLOCK_UNINIT` is set for a given block group, various parts of the kernel and e2fsprogs code pretends that the block bitmap contains zeros (i.e. all blocks in the group are free). However, it is not necessarily the case that no blocks are in use -- if `meta_bg` is set, the bitmaps and group descriptor live inside the group. Unfortunately, ext2fs_test_block_bitmap2() will return '0' for those locations, which produces confusing debugfs output.

### Inode Table

Inode tables are statically allocated at mkfs time. Each block group descriptor points to the start of the table, and the superblock records the number of inodes per group. See the section on inodes for more information.

### Multiple Mount Protection

Multiple mount protection (MMP) is a feature that protects the filesystem against multiple hosts trying to use the filesystem simultaneously. When a filesystem is opened (for mounting, or fsck, etc.), the MMP code running on the node (call it node A) checks a sequence number. If the sequence number is EXT4_MMP_SEQ_CLEAN, the open continues. If the sequence number is EXT4_MMP_SEQ_FSCK, then fsck is (hopefully) running, and open fails immediately. Otherwise, the open code will wait for twice the specified MMP check interval and check the sequence number again. If the sequence number has changed, then the filesystem is active on another machine and the open fails. If the MMP code passes all of those checks, a new MMP sequence number is generated and written to the MMP block, and the mount proceeds.

While the filesystem is live, the kernel sets up a timer to re-check the MMP block at the specified MMP check interval. To perform the re-check, the MMP sequence number is re-read; if it does not match the in-memory MMP sequence number, then another node (node B) has mounted the filesystem, and node A remounts the filesystem read-only. If the sequence numbers match, the sequence number is incremented both in memory and on disk, and the re-check is complete.

The hostname and device filename are written into the MMP block whenever an open operation succeeds. The MMP code does not use these values; they are provided purely for informational purposes.

The checksum is calculated against the FS UUID and the MMP structure. The MMP structure (`struct mmp_struct`) is as follows:

| Offset | Type | Name | Description |
|---|---|---|---|
| 0x0 | __le32 | mmp_magic | Magic number for MMP, 0x004D4D50 (“MMP”). |
| 0x4 | __le32 | mmp_seq | Sequence number, updated periodically. |
| 0x8 | __le64 | mmp_time | Time that the MMP block was last updated. |
| 0x10 | char[64] | mmp_nodename | Hostname of the node that opened the filesystem. |
| 0x50 | char[32] | mmp_bdevname | Block device name of the filesystem. |
| 0x70 | __le16 | mmp_check_interval | The MMP re-check interval, in seconds. |
| 0x72 | __le16 | mmp_pad1 | Zero. |
| 0x74 | __le32[226] | mmp_pad2 | Zero. |
| 0x3FC | __le32 | mmp_checksum | Checksum of the MMP block. |

### Journal (jbd2)

Introduced in ext3, the ext4 filesystem employs a journal to protect the filesystem against metadata inconsistencies in the case of a system crash. Up to 10,240,000 file system blocks (see man mke2fs(8) for more details on journal size limits) can be reserved inside the filesystem as a place to land “important” data writes on-disk as quickly as possible. Once the important data transaction is fully written to the disk and flushed from the disk write cache, a record of the data being committed is also written to the journal. At some later point in time, the journal code writes the transactions to their final locations on disk (this could involve a lot of seeking or a lot of small read-write-erases) before erasing the commit record. Should the system crash during the second slow write, the journal can be replayed all the way to the latest commit record, guaranteeing the atomicity of whatever gets written through the journal to the disk. The effect of this is to guarantee that the filesystem does not become stuck midway through a metadata update.

For performance reasons, ext4 by default only writes filesystem metadata through the journal. This means that file data blocks are /not/ guaranteed to be in any consistent state after a crash. If this default guarantee level (`data=ordered`) is not satisfactory, there is a mount option to control journal behavior. If `data=journal`, all data and metadata are written to disk through the journal. This is slower but safest. If `data=writeback`, dirty data blocks are not flushed to the disk before the metadata are written to disk through the journal.

In case of `data=ordered` mode, Ext4 also supports fast commits which help reduce commit latency significantly. The default `data=ordered` mode works by logging metadata blocks to the journal. In fast commit mode, Ext4 only stores the minimal delta needed to recreate the affected metadata in fast commit space that is shared with JBD2. Once the fast commit area fills in or if fast commit is not possible or if JBD2 commit timer goes off, Ext4 performs a traditional full commit. A full commit invalidates all the fast commits that happened before it and thus it makes the fast commit area empty for further fast commits. This feature needs to be enabled at mkfs time.

The journal inode is typically inode 8. The first 68 bytes of the journal inode are replicated in the ext4 superblock. The journal itself is normal (but hidden) file within the filesystem. The file usually consumes an entire block group, though mke2fs tries to put it in the middle of the disk.

All fields in jbd2 are written to disk in big-endian order. This is the opposite of ext4.

NOTE: Both ext4 and ocfs2 use jbd2.

The maximum size of a journal embedded in an ext4 filesystem is 2^32 blocks. jbd2 itself does not seem to care.

### Layout

Generally speaking, the journal has this format:

| Superblock | descriptor_block (data_blocks or revocation_block) [more data or revocations] commmit_block | [more transactions...] |
|---|---|---|
| | One transaction | |

Notice that a transaction begins with either a descriptor and some data, or a block revocation list. A finished transaction always ends with a commit. If there is no commit record (or the checksums don't match), the transaction will be discarded during replay.

### External Journal

Optionally, an ext4 filesystem can be created with an external journal device (as opposed to an internal journal, which uses a reserved inode). In this case, on the filesystem device, `s_journal_inum` should be zero and `s_journal_uuid` should be set. On the journal device there will be an ext4 super block in the usual place, with a matching UUID. The journal superblock will be in the next full block after the superblock.

| 1024 bytes of padding | ext4 Superblock | Journal Superblock | descriptor_block (data_blocks or revocation_block) [more data or revocations] commmit_block | [more transactions...] |
|---|---|---|---|---|
| | | | One transaction | |

### Block Header

Every block in the journal starts with a common 12-byte header `struct journal_header_s`:

| Offset | Type | Name | Description |
|---|---|---|---|
| 0x0 | __be32 | h_magic | jbd2 magic number, 0xC03B3998. |
| 0x4 | __be32 | h_blocktype | Description of what this block contains. See the *jbd2_blocktype* table below. |
| 0x8 | __be32 | h_sequence | The transaction ID that goes with this block. |

The journal block type can be any one of:

| Value | Description |
|---|---|
| 1 | Descriptor. This block precedes a series of data blocks that were written through the journal during a transaction. |
| 2 | Block commit record. This block signifies the completion of a transaction. |
| 3 | Journal superblock, v1. |
| 4 | Journal superblock, v2. |
| 5 | Block revocation records. This speeds up recovery by enabling the journal to skip writing blocks that were subsequently rewritten. |

### Super Block

The super block for the journal is much simpler as compared to ext4's. The key data kept within are size of the journal, and where to find the start of the log of transactions.

The journal superblock is recorded as `struct journal_superblock_s`, which is 1024 bytes long:

| Offset | Type | Name | Description |
|---|---|---|---|
| | | | Static information describing the journal. |
| 0x0 | journal_header_t (12 bytes) | s_header | Common header identifying this as a superblock. |
| 0xC | __be32 | s_blocksize | Journal device block size. |
| 0x10 | __be32 | s_maxlen | Total number of blocks in this journal. |
| 0x14 | __be32 | s_first | First block of log information. |
| | | | Dynamic information describing the current state of the log. |
| 0x18 | __be32 | s_sequence | First commit ID expected in log. |
| 0x1C | __be32 | s_start | Block number of the start of log. Contrary to the comments, this field being zero does not imply that the journal is clean! |
| 0x20 | __be32 | s_errno | Error value, as set by *jbd2_journal_abort()*. |
| | | | The remaining fields are only valid in a v2 superblock. |
| 0x24 | __be32 | s_feature_compat; | Compatible feature set. See the table *jbd2_compat* below. |
| 0x28 | __be32 | s_feature_incompat | Incompatible feature set. See the table *jbd2_incompat* below. |
| 0x2C | __be32 | s_feature_ro_compat | Read-only compatible feature set. There aren't any of these currently. |
| 0x30 | __u8 | s_uuid[16] | 128-bit uuid for journal. This is compared against the copy in the ext4 super block at mount time. |
| 0x40 | __be32 | s_nr_users | Number of file systems sharing this journal. |
| 0x44 | __be32 | s_dynsuper | Location of dynamic super block copy. (Not used?) |
| 0x48 | __be32 | s_max_transaction | Limit of journal blocks per transaction. (Not used?) |
| 0x4C | __be32 | s_max_trans_data | Limit of data blocks per transaction. (Not used?) |
| 0x50 | __u8 | s_checksum_type | Checksum algorithm used for the journal. See *jbd2_checksum_type* for more info. |
| 0x51 | __u8[3] | s_padding2 | |
| 0x54 | __be32 | s_num_fc_blocks | Number of fast commit blocks in the journal. |
| 0x58 | __be32 | s_head | Block number of the head (first unused block) of the journal, only up-to-date when the journal is empty. |
| 0x5C | __u32 | s_padding[40] | |
| 0xFC | __be32 | s_checksum | Checksum of the entire superblock, with this field set to zero. |
| 0x100 | __u8 | s_users[16*48] | ids of all file systems sharing the log. e2fsprogs/Linux don't allow shared external journals, but I imagine Lustre (or ocfs2?), which use the jbd2 code, might. |

The journal compat features are any combination of the following:

| Value | Description |
|---|---|
| 0x1 | Journal maintains checksums on the data blocks. (JBD2_FEATURE_COMPAT_CHECKSUM) |

The journal incompat features are any combination of the following:

| Value | Description |
|---|---|
| 0x1 | Journal has block revocation records. (JBD2_FEATURE_INCOMPAT_REVOKE) |
| 0x2 | Journal can deal with 64-bit block numbers. (JBD2_FEATURE_INCOMPAT_64BIT) |
| 0x4 | Journal commits asynchronously. (JBD2_FEATURE_INCOMPAT_ASYNC_COMMIT |
| 0x8 | This journal uses v2 of the checksum on-disk format. Each journal metadata block gets its own checksum, and the block tags in the descriptor table contain checksums for each of the data blocks in the journal. (JBD2_FEATURE_INCOMPAT_CSUM_V2) |
| 0x10 | This journal uses v3 of the checksum on-disk format. This is the same as v2, but the journal block tag size is fixed regardless of the size of block numbers. (JBD2_FEATURE_INCOMPAT_CSUM_V3) |
| 0x20 | Journal has fast commit blocks. (JBD2_FEATURE_INCOMPAT_FAST_COMMIT) |

Journal checksum type codes are one of the following. crc32 or crc32c are the most likely choices.

| Value | Description |
|---|---|
| 1 | CRC32 |
| 2 | MD5 |
| 3 | SHA1 |
| 4 | CRC32C |

### Descriptor Block

The descriptor block contains an array of journal block tags that describe the final locations of the data blocks that follow in the journal. Descriptor blocks are open-coded instead of being completely described by a data structure, but here is the block structure anyway. Descriptor blocks consume at least 36 bytes, but use a full block:

| Offset | Type | Name | Descriptor |
|---|---|---|---|
| 0x0 | journal_header_t | (open coded) | Common block header. |
| 0xC | struct journal_block_tag_s | open coded array[] | Enough tags either to fill up the block or to describe all the data blocks that follow this descriptor block. |

Journal block tags have any of the following formats, depending on which journal feature and block tag flags are set.

If JBD2_FEATURE_INCOMPAT_CSUM_V3 is set, the journal block tag is defined as `struct journal_block_tag3_s`, which looks like the following. The size is 16 or 32 bytes.

| Offset | Type | Name | Descriptor |
|---|---|---|---|
| 0x0 | __be32 | t_blocknr | Lower 32-bits of the location of where the corresponding data block should end up on disk. |
| 0x4 | __be32 | t_flags | Flags that go with the descriptor. See the table *jbd2_tag_flags* for more info. |
| 0x8 | __be32 | t_blocknr_high | Upper 32-bits of the location of where the corresponding data block should end up on disk. This is zero if JBD2_FEATURE_INCOMPAT_64BIT is not enabled. |
| 0xC | __be32 | t_checksum | Checksum of the journal UUID, the sequence number, and the data block. |
| | | | This field appears to be open coded. It always comes at the end of the tag, after t_checksum. This field is not present if the "same UUID" flag is set. |
| 0x8 or 0xC | char | uuid[16] | A UUID to go with this tag. This field appears to be copied from the `j_uuid` field in `struct journal_s`, but only tune2fs touches that field. |

The journal tag flags are any combination of the following:

| Value | Description |
|---|---|
| 0x1 | On-disk block is escaped. The first four bytes of the data block just happened to match the jbd2 magic number. |
| 0x2 | This block has the same UUID as previous, therefore the UUID field is omitted. |
| 0x4 | The data block was deleted by the transaction. (Not used?) |
| 0x8 | This is the last tag in this descriptor block. |

If JBD2_FEATURE_INCOMPAT_CSUM_V3 is NOT set, the journal block tag is defined as `struct journal_block_tag_s`, which looks like the following. The size is 8, 12, 24, or 28 bytes:

| Offset | Type | Name | Descriptor |
|---|---|---|---|
| 0x0 | __be32 | t_blocknr | Lower 32-bits of the location of where the corresponding data block should end up on disk. |
| 0x4 | __be16 | t_checksum | Checksum of the journal UUID, the sequence number, and the data block. Note that only the lower 16 bits are stored. |
| 0x6 | __be16 | t_flags | Flags that go with the descriptor. See the table *jbd2_tag_flags* for more info. |
| | | | This next field is only present if the super block indicates support for 64-bit block numbers. |
| 0x8 | __be32 | t_blocknr_high | Upper 32-bits of the location of where the corresponding data block should end up on disk. |
| | | | This field appears to be open coded. It always comes at the end of the tag, after t_flags or t_blocknr_high. This field is not present if the "same UUID" flag is set. |
| 0x8 or 0xC | char | uuid[16] | A UUID to go with this tag. This field appears to be copied from the `j_uuid` field in `struct journal_s`, but only tune2fs touches that field. |

If JBD2_FEATURE_INCOMPAT_CSUM_V2 or JBD2_FEATURE_INCOMPAT_CSUM_V3 are set, the end of the block is a `struct jbd2_journal_block_tail`, which looks like this:

| Offset | Type | Name | Descriptor |
|---|---|---|---|
| 0x0 | __be32 | t_checksum | Checksum of the journal UUID + the descriptor block, with this field set to zero. |

### Data Block

In general, the data blocks being written to disk through the journal are written verbatim into the journal file after the descriptor block. However, if the first four bytes of the block match the jbd2 magic number then those four bytes are replaced with zeroes and the "escaped" flag is set in the descriptor block tag.

### Revocation Block

A revocation block is used to prevent replay of a block in an earlier transaction. This is used to mark blocks that were journalled at one time but are no longer journalled. Typically this happens if a metadata block is freed and re-allocated as a file data block; in this case, a journal replay after the file block was written to disk will cause corruption.

**NOTE**: This mechanism is NOT used to express "this journal block is superseded by this other journal block", as the author (djwong) mistakenly thought. Any block being added to a transaction will cause the removal of all existing revocation records for that block.

Revocation blocks are described in `struct jbd2_journal_revoke_header_s`, are at least 16 bytes in length, but use a full block:

| Offset | Type | Name | Description |
|---|---|---|---|
| 0x0 | jour-nal_header_t | r_header | Common block header. |
| 0xC | __be32 | r_count | Number of bytes used in this block. |
| 0x10 | __be32 or __be64 | blocks[0] | Blocks to revoke. |

After r_count is a linear array of block numbers that are effectively revoked by this transaction. The size of each block number is 8 bytes if the superblock advertises 64-bit block number support, or 4 bytes otherwise.

If JBD2_FEATURE_INCOMPAT_CSUM_V2 or JBD2_FEATURE_INCOMPAT_CSUM_V3 are set, the end of the revocation block is a `struct jbd2_journal_revoke_tail`, which has this format:

| Offset | Type | Name | Description |
|---|---|---|---|
| 0x0 | __be32 | r_checksum | Checksum of the journal UUID + revocation block |

### Commit Block

The commit block is a sentry that indicates that a transaction has been completely written to the journal. Once this commit block reaches the journal, the data stored with this transaction can be written to their final locations on disk.

The commit block is described by `struct commit_header`, which is 32 bytes long (but uses a full block):

| Offset | Type | Name | Descriptor |
|---|---|---|---|
| 0x0 | journal_header_s | (open coded) | Common block header. |
| 0xC | unsigned char | h_chksum_type | The type of checksum to use to verify the integrity of the data blocks in the transaction. See *jbd2_checksum_type* for more info. |
| 0xD | unsigned char | h_chksum_size | The number of bytes used by the checksum. Most likely 4. |
| 0xE | unsigned char | h_padding[2] | |
| 0x10 | __be32 | h_chksum[JBD2_CHECKSUM_BYTES] | 32 bytes of space to store checksums. If JBD2_FEATURE_INCOMPAT_CSUM_V2 or JBD2_FEATURE_INCOMPAT_CSUM_V3 are set, the first `__be32` is the checksum of the journal UUID and the entire commit block, with this field zeroed. If JBD2_FEATURE_COMPAT_CHECKSUM is set, the first `__be32` is the crc32 of all the blocks already written to the transaction. |
| 0x30 | __be64 | h_commit_sec | The time that the transaction was committed, in seconds since the epoch. |
| 0x38 | __be32 | h_commit_nsec | Nanoseconds component of the above timestamp. |

### Fast commits

Fast commit area is organized as a log of tag length values. Each TLV has a `struct ext4_fc_tl` in the beginning which stores the tag and the length of the entire field. It is followed by variable length tag specific value. Here is the list of supported tags and their meanings:

| Tag | Meaning | Value struct | Description |
|---|---|---|---|
| EXT4_FC_TAG_HEAD | Fast commit area header | `struct ext4_fc_head` | Stores the TID of the transaction after which these fast commits should be applied. |
| EXT4_FC_TAG_ADD_RANGE | Add extent to inode | `struct ext4_fc_add_range` | Stores the inode number and extent to be added in this inode |
| EXT4_FC_TAG_DEL_RANGE | Remove logical offsets to inode | `struct ext4_fc_del_range` | Stores the inode number and the logical offset range that needs to be removed |
| EXT4_FC_TAG_CREAT | Create directory entry for a newly created file | `struct ext4_fc_dentry_info` | Stores the parent inode number, inode number and directory entry of the newly created file |
| EXT4_FC_TAG_LINK | Link directory entry to an inode | `struct ext4_fc_dentry_info` | Stores the parent inode number, inode number and directory entry |
| EXT4_FC_TAG_UNLINK | Unlink directory entry of an inode | `struct ext4_fc_dentry_info` | Stores the parent inode number, inode number and directory entry |
| EXT4_FC_TAG_PAD | Padding (unused area) | None | Unused bytes in the fast commit area. |
| EXT4_FC_TAG_TAIL | Mark the end of a fast commit | `struct ext4_fc_tail` | Stores the TID of the commit, CRC of the fast commit of which this tag represents the end of |

### Fast Commit Replay Idempotence

Fast commits tags are idempotent in nature provided the recovery code follows certain rules. The guiding principle that the commit path follows while committing is that it stores the result of a particular operation instead of storing the procedure.

Let's consider this rename operation: 'mv /a /b'. Let's assume dirent '/a' was associated with inode 10. During fast commit, instead of storing this operation as a procedure ”rename a to b”, we store the resulting file system state as a ”series” of outcomes:

- Link dirent b to inode 10
- Unlink dirent a
- Inode 10 with valid refcount

Now when recovery code runs, it needs ”enforce” this state on the file system. This is what guarantees idempotence of fast commit replay.

Let's take an example of a procedure that is not idempotent and see how fast commits make it idempotent. Consider following sequence of operations:

1) rm A
2) mv B A
3) read A

If we store this sequence of operations as is then the replay is not idempotent. Let's say while in replay, we crash after (2). During the second replay, file A (which was actually created as a result of ”mv B A” operation) would get deleted. Thus, file named A would be absent when we try to read A. So, this sequence of operations is not idempotent. However, as mentioned above,

instead of storing the procedure fast commits store the outcome of each procedure. Thus the fast commit log for above procedure would be as follows:

(Let's assume dirent A was linked to inode 10 and dirent B was linked to inode 11 before the replay)

1) Unlink A
2) Link A to inode 11
3) Unlink B
4) Inode 11

If we crash after (3) we will have file A linked to inode 11. During the second replay, we will remove file A (inode 11). But we will create it back and make it point to inode 11. We won't find B, so we'll just skip that step. At this point, the refcount for inode 11 is not reliable, but that gets fixed by the replay of last inode 11 tag. Thus, by converting a non-idempotent procedure into a series of idempotent outcomes, fast commits ensured idempotence during the replay.

### Journal Checkpoint

Checkpointing the journal ensures all transactions and their associated buffers are submitted to the disk. In-progress transactions are waited upon and included in the checkpoint. Checkpointing is used internally during critical updates to the filesystem including journal recovery, filesystem resizing, and freeing of the journal_t structure.

A journal checkpoint can be triggered from userspace via the ioctl EXT4_IOC_CHECKPOINT. This ioctl takes a single, u64 argument for flags. Currently, three flags are supported. First, EXT4_IOC_CHECKPOINT_FLAG_DRY_RUN can be used to verify input to the ioctl. It returns error if there is any invalid input, otherwise it returns success without performing any checkpointing. This can be used to check whether the ioctl exists on a system and to verify there are no issues with arguments or flags. The other two flags are EXT4_IOC_CHECKPOINT_FLAG_DISCARD and EXT4_IOC_CHECKPOINT_FLAG_ZEROOUT. These flags cause the journal blocks to be discarded or zero-filled, respectively, after the journal checkpoint is complete. EXT4_IOC_CHECKPOINT_FLAG_DISCARD and EXT4_IOC_CHECKPOINT_FLAG_ZEROOUT cannot both be set. The ioctl may be useful when snapshotting a system or for complying with content deletion SLOs.

### Orphan file

In unix there can inodes that are unlinked from directory hierarchy but that are still alive because they are open. In case of crash the filesystem has to clean up these inodes as otherwise they (and the blocks referenced from them) would leak. Similarly if we truncate or extend the file, we need not be able to perform the operation in a single journalling transaction. In such case we track the inode as orphan so that in case of crash extra blocks allocated to the file get truncated.

Traditionally ext4 tracks orphan inodes in a form of single linked list where superblock contains the inode number of the last orphan inode (s_last_orphan field) and then each inode contains inode number of the previously orphaned inode (we overload i_dtime inode field for this). However this filesystem global single linked list is a scalability bottleneck for

workloads that result in heavy creation of orphan inodes. When orphan file feature (COMPAT_ORPHAN_FILE) is enabled, the filesystem has a special inode (referenced from the superblock through s_orphan_file_inum) with several blocks. Each of these blocks has a structure:

| Offset | Type | Name | Description |
|---|---|---|---|
| 0x0 | Array of __le32 entries | Orphan inode entries | Each __le32 entry is either empty (0) or it contains inode number of an orphan inode. |
| blocksize-8 | __le32 | ob_magic | Magic value stored in orphan block tail (0x0b10ca04) |
| blocksize-4 | __le32 | ob_checksum | Checksum of the orphan block. |

When a filesystem with orphan file feature is writeably mounted, we set RO_COMPAT_ORPHAN_PRESENT feature in the superblock to indicate there may be valid orphan entries. In case we see this feature when mounting the filesystem, we read the whole orphan file and process all orphan inodes found there as usual. When cleanly unmounting the filesystem we remove the RO_COMPAT_ORPHAN_PRESENT feature to avoid unnecessary scanning of the orphan file and also make the filesystem fully compatible with older kernels.

### 3.22.4 Dynamic Structures

Dynamic metadata are created on the fly when files and blocks are allocated to files.

#### Index Nodes

In a regular UNIX filesystem, the inode stores all the metadata pertaining to the file (time stamps, block maps, extended attributes, etc), not the directory entry. To find the information associated with a file, one must traverse the directory files to find the directory entry associated with a file, then load the inode to find the metadata for that file. ext4 appears to cheat (for performance reasons) a little bit by storing a copy of the file type (normally stored in the inode) in the directory entry. (Compare all this to FAT, which stores all the file information directly in the directory entry, but does not support hard links and is in general more seek-happy than ext4 due to its simpler block allocator and extensive use of linked lists.)

The inode table is a linear array of `struct ext4_inode`. The table is sized to have enough blocks to store at least `sb.s_inode_size * sb.s_inodes_per_group` bytes. The number of the block group containing an inode can be calculated as `(inode_number - 1) / sb.s_inodes_per_group`, and the offset into the group's table is `(inode_number - 1) % sb.s_inodes_per_group`. There is no inode 0.

The inode checksum is calculated against the FS UUID, the inode number, and the inode structure itself.

The inode table entry is laid out in `struct ext4_inode`.

| Offset | Size | Name | Description |
|---|---|---|---|
| 0x0 | __le16 | i_mode | File mode. See the table *i_mode* below. |
| 0x2 | __le16 | i_uid | Lower 16-bits of Owner UID. |
| 0x4 | __le32 | i_size_lo | Lower 32-bits of size in bytes. |

continues on next page

Table 2 – continued from previous page

| Offset | Size | Name | Description |
|---|---|---|---|
| 0x8 | __le32 | i_atime | Last access time, in seconds since the epoch. However, if the EA_INODE inode flag is set, this inode stores an extended attribute value and this field contains the checksum of the value. |
| 0xC | __le32 | i_ctime | Last inode change time, in seconds since the epoch. However, if the EA_INODE inode flag is set, this inode stores an extended attribute value and this field contains the lower 32 bits of the attribute value's reference count. |
| 0x10 | __le32 | i_mtime | Last data modification time, in seconds since the epoch. However, if the EA_INODE inode flag is set, this inode stores an extended attribute value and this field contains the number of the inode that owns the extended attribute. |
| 0x14 | __le32 | i_dtime | Deletion Time, in seconds since the epoch. |
| 0x18 | __le16 | i_gid | Lower 16-bits of GID. |
| 0x1A | __le16 | i_links_count | Hard link count. Normally, ext4 does not permit an inode to have more than 65,000 hard links. This applies to files as well as directories, which means that there cannot be more than 64,998 subdirectories in a directory (each subdirectory's '..' entry counts as a hard link, as does the '.' entry in the directory itself). With the DIR_NLINK feature enabled, ext4 supports more than 64,998 subdirectories by setting this field to 1 to indicate that the number of hard links is not known. |
| 0x1C | __le32 | i_blocks_lo | Lower 32-bits of "block" count. If the huge_file feature flag is not set on the filesystem, the file consumes `i_blocks_lo` 512-byte blocks on disk. If huge_file is set and EXT4_HUGE_FILE_FL is NOT set in `inode.i_flags`, then the file consumes `i_blocks_lo + (i_blocks_hi << 32)` 512-byte blocks on disk. If huge_file is set and EXT4_HUGE_FILE_FL IS set in `inode.i_flags`, then this file consumes `(i_blocks_lo + i_blocks_hi << 32)` filesystem blocks on disk. |
| 0x20 | __le32 | i_flags | Inode flags. See the table *i_flags* below. |
| 0x24 | 4 bytes | i_osd1 | See the table *i_osd1* for more details. |
| 0x28 | 60 bytes | i_block[EXT4_N_BLOCKS=15] | Block map or extent tree. See the section "The Contents of inode.i_block". |
| 0x64 | __le32 | i_generation | File version (for NFS). |

continues on next page

Table 2 – continued from previous page

| Offset | Size | Name | Description |
|---|---|---|---|
| 0x68 | __le32 | i_file_acl_lo | Lower 32-bits of extended attribute block. ACLs are of course one of many possible extended attributes; I think the name of this field is a result of the first use of extended attributes being for ACLs. |
| 0x6C | __le32 | i_size_high / i_dir_acl | Upper 32-bits of file/directory size. In ext2/3 this field was named i_dir_acl, though it was usually set to zero and never used. |
| 0x70 | __le32 | i_obso_faddr | (Obsolete) fragment address. |
| 0x74 | 12 bytes | i_osd2 | See the table *i_osd2* for more details. |
| 0x80 | __le16 | i_extra_isize | Size of this inode - 128. Alternately, the size of the extended inode fields beyond the original ext2 inode, including this field. |
| 0x82 | __le16 | i_checksum_hi | Upper 16-bits of the inode checksum. |
| 0x84 | __le32 | i_ctime_extra | Extra change time bits. This provides sub-second precision. See Inode Timestamps section. |
| 0x88 | __le32 | i_mtime_extra | Extra modification time bits. This provides sub-second precision. |
| 0x8C | __le32 | i_atime_extra | Extra access time bits. This provides sub-second precision. |
| 0x90 | __le32 | i_crtime | File creation time, in seconds since the epoch. |
| 0x94 | __le32 | i_crtime_extra | Extra file creation time bits. This provides sub-second precision. |
| 0x98 | __le32 | i_version_hi | Upper 32-bits for version number. |
| 0x9C | __le32 | i_projid | Project ID. |

The `i_mode` value is a combination of the following flags:

| Value | Description |
|---|---|
| 0x1 | S_IXOTH (Others may execute) |
| 0x2 | S_IWOTH (Others may write) |
| 0x4 | S_IROTH (Others may read) |
| 0x8 | S_IXGRP (Group members may execute) |
| 0x10 | S_IWGRP (Group members may write) |
| 0x20 | S_IRGRP (Group members may read) |
| 0x40 | S_IXUSR (Owner may execute) |
| 0x80 | S_IWUSR (Owner may write) |
| 0x100 | S_IRUSR (Owner may read) |
| 0x200 | S_ISVTX (Sticky bit) |
| 0x400 | S_ISGID (Set GID) |
| 0x800 | S_ISUID (Set UID) |
| | These are mutually-exclusive file types: |
| 0x1000 | S_IFIFO (FIFO) |
| 0x2000 | S_IFCHR (Character device) |
| 0x4000 | S_IFDIR (Directory) |
| 0x6000 | S_IFBLK (Block device) |
| 0x8000 | S_IFREG (Regular file) |
| 0xA000 | S_IFLNK (Symbolic link) |
| 0xC000 | S_IFSOCK (Socket) |

The `i_flags` field is a combination of these values:

| Value | Description |
|---|---|
| 0x1 | This file requires secure deletion (EXT4_SECRM_FL). (not implemented) |
| 0x2 | This file should be preserved, should undeletion be desired (EXT4_UNRM_FL). (not implemented) |
| 0x4 | File is compressed (EXT4_COMPR_FL). (not really implemented) |
| 0x8 | All writes to the file must be synchronous (EXT4_SYNC_FL). |
| 0x10 | File is immutable (EXT4_IMMUTABLE_FL). |
| 0x20 | File can only be appended (EXT4_APPEND_FL). |
| 0x40 | The dump(1) utility should not dump this file (EXT4_NODUMP_FL). |
| 0x80 | Do not update access time (EXT4_NOATIME_FL). |
| 0x100 | Dirty compressed file (EXT4_DIRTY_FL). (not used) |
| 0x200 | File has one or more compressed clusters (EXT4_COMPRBLK_FL). (not used) |
| 0x400 | Do not compress file (EXT4_NOCOMPR_FL). (not used) |
| 0x800 | Encrypted inode (EXT4_ENCRYPT_FL). This bit value previously was EXT4_ECOMPR_FL (compression error), which was never used. |
| 0x1000 | Directory has hashed indexes (EXT4_INDEX_FL). |
| 0x2000 | AFS magic directory (EXT4_IMAGIC_FL). |
| 0x4000 | File data must always be written through the journal (EXT4_JOURNAL_DATA_FL). |
| 0x8000 | File tail should not be merged (EXT4_NOTAIL_FL). (not used by ext4) |
| 0x10000 | All directory entry data should be written synchronously (see `dirsync`) (EXT4_DIRSYNC_FL). |
| 0x20000 | Top of directory hierarchy (EXT4_TOPDIR_FL). |

continues on next page

Table 3 – continued from previous page

| Value | Description |
|---|---|
| 0x40000 | This is a huge file (EXT4_HUGE_FILE_FL). |
| 0x80000 | Inode uses extents (EXT4_EXTENTS_FL). |
| 0x100000 | Verity protected file (EXT4_VERITY_FL). |
| 0x200000 | Inode stores a large extended attribute value in its data blocks (EXT4_EA_INODE_FL). |
| 0x400000 | This file has blocks allocated past EOF (EXT4_EOFBLOCKS_FL). (deprecated) |
| 0x01000000 | Inode is a snapshot (EXT4_SNAPFILE_FL). (not in mainline) |
| 0x04000000 | Snapshot is being deleted (EXT4_SNAPFILE_DELETED_FL). (not in mainline) |
| 0x08000000 | Snapshot shrink has completed (EXT4_SNAPFILE_SHRUNK_FL). (not in mainline) |
| 0x10000000 | Inode has inline data (EXT4_INLINE_DATA_FL). |
| 0x20000000 | Create children with the same project ID (EXT4_PROJINHERIT_FL). |
| 0x80000000 | Reserved for ext4 library (EXT4_RESERVED_FL). |
| | Aggregate flags: |
| 0x705BDFFF | User-visible flags. |
| 0x604BC0FF | User-modifiable flags. Note that while EXT4_JOURNAL_DATA_FL and EXT4_EXTENTS_FL can be set with setattr, they are not in the kernel's EXT4_FL_USER_MODIFIABLE mask, since it needs to handle the setting of these flags in a special manner and they are masked out of the set of flags that are saved directly to i_flags. |

The `osd1` field has multiple meanings depending on the creator:

Linux:

| Offset | Size | Name | Description |
|---|---|---|---|
| 0x0 | __le32 | l_i_version | Inode version. However, if the EA_INODE inode flag is set, this inode stores an extended attribute value and this field contains the upper 32 bits of the attribute value's reference count. |

Hurd:

| Offset | Size | Name | Description |
|---|---|---|---|
| 0x0 | __le32 | h_i_translator | ?? |

Masix:

| Offset | Size | Name | Description |
|---|---|---|---|
| 0x0 | __le32 | m_i_reserved | ?? |

The `osd2` field has multiple meanings depending on the filesystem creator:

Linux:

| Offset | Size | Name | Description |
|---|---|---|---|
| 0x0 | __le16 | l_i_blocks_high | Upper 16-bits of the block count. Please see the note attached to i_blocks_lo. |
| 0x2 | __le16 | l_i_file_acl_high | Upper 16-bits of the extended attribute block (historically, the file ACL location). See the Extended Attributes section below. |
| 0x4 | __le16 | l_i_uid_high | Upper 16-bits of the Owner UID. |
| 0x6 | __le16 | l_i_gid_high | Upper 16-bits of the GID. |
| 0x8 | __le16 | l_i_checksum_lo | Lower 16-bits of the inode checksum. |
| 0xA | __le16 | l_i_reserved | Unused. |

Hurd:

| Offset | Size | Name | Description |
|---|---|---|---|
| 0x0 | __le16 | h_i_reserved1 | ?? |
| 0x2 | __u16 | h_i_mode_high | Upper 16-bits of the file mode. |
| 0x4 | __le16 | h_i_uid_high | Upper 16-bits of the Owner UID. |
| 0x6 | __le16 | h_i_gid_high | Upper 16-bits of the GID. |
| 0x8 | __u32 | h_i_author | Author code? |

Masix:

| Offset | Size | Name | Description |
|---|---|---|---|
| 0x0 | __le16 | h_i_reserved1 | ?? |
| 0x2 | __u16 | m_i_file_acl_high | Upper 16-bits of the extended attribute block (historically, the file ACL location). |
| 0x4 | __u32 | m_i_reserved2[2] | ?? |

### Inode Size

In ext2 and ext3, the inode structure size was fixed at 128 bytes (`EXT2_GOOD_OLD_INODE_SIZE`) and each inode had a disk record size of 128 bytes. Starting with ext4, it is possible to allocate a larger on-disk inode at format time for all inodes in the filesystem to provide space beyond the end of the original ext2 inode. The on-disk inode record size is recorded in the superblock as `s_inode_size`. The number of bytes actually used by struct ext4_inode beyond the original 128-byte ext2 inode is recorded in the `i_extra_isize` field for each inode, which allows struct ext4_inode to grow for a new kernel without having to upgrade all of the on-disk inodes. Access to fields beyond EXT2_GOOD_OLD_INODE_SIZE should be verified to be within `i_extra_isize`. By default, ext4 inode records are 256 bytes, and (as of August 2019) the inode structure is 160 bytes (`i_extra_isize = 32`). The extra space between the end of the inode structure and the end of the inode record can be used to store extended attributes. Each inode record can be as large as the filesystem block size, though this is not terribly efficient.

### Finding an Inode

Each block group contains `sb->s_inodes_per_group` inodes. Because inode 0 is defined not to exist, this formula can be used to find the block group that an inode lives in: `bg = (inode_num - 1) / sb->s_inodes_per_group`. The particular inode can be found within the block group's inode table at `index = (inode_num - 1) % sb->s_inodes_per_group`. To get the byte address within the inode table, use `offset = index * sb->s_inode_size`.

### Inode Timestamps

Four timestamps are recorded in the lower 128 bytes of the inode structure -- inode change time (ctime), access time (atime), data modification time (mtime), and deletion time (dtime). The four fields are 32-bit signed integers that represent seconds since the Unix epoch (1970-01-01 00:00:00 GMT), which means that the fields will overflow in January 2038. If the filesystem does not have orphan_file feature, inodes that are not linked from any directory but are still open (orphan inodes) have the dtime field overloaded for use with the orphan list. The superblock field `s_last_orphan` points to the first inode in the orphan list; dtime is then the number of the next orphaned inode, or zero if there are no more orphans.

If the inode structure size `sb->s_inode_size` is larger than 128 bytes and the `i_inode_extra` field is large enough to encompass the respective `i_[cma]time_extra` field, the ctime, atime, and mtime inode fields are widened to 64 bits. Within this "extra" 32-bit field, the lower two bits are used to extend the 32-bit seconds field to be 34 bit wide; the upper 30 bits are used to provide nanosecond timestamp accuracy. Therefore, timestamps should not overflow until May 2446. dtime was not widened. There is also a fifth timestamp to record inode creation time (crtime); this field is 64-bits wide and decoded in the same manner as 64-bit [cma]time. Neither crtime nor dtime are accessible through the regular stat() interface, though debugfs will report them.

We use the 32-bit signed time value plus (2^32 * (extra epoch bits)). In other words:

| Extra epoch bits | MSB of 32-bit time | Adjustment for signed 32-bit to 64-bit tv_sec | Decoded 64-bit tv_sec | valid time range |
|---|---|---|---|---|
| 0 0 | 1 | 0 | `-0x80000000` - `-0x00000001` | 1901-12-13 to 1969-12-31 |
| 0 0 | 0 | 0 | `0x000000000` - `0x07fffffff` | 1970-01-01 to 2038-01-19 |
| 0 1 | 1 | 0x100000000 | `0x080000000` - `0x0ffffffff` | 2038-01-19 to 2106-02-07 |
| 0 1 | 0 | 0x100000000 | `0x100000000` - `0x17fffffff` | 2106-02-07 to 2174-02-25 |
| 1 0 | 1 | 0x200000000 | `0x180000000` - `0x1ffffffff` | 2174-02-25 to 2242-03-16 |
| 1 0 | 0 | 0x200000000 | `0x200000000` - `0x27fffffff` | 2242-03-16 to 2310-04-04 |
| 1 1 | 1 | 0x300000000 | `0x280000000` - `0x2ffffffff` | 2310-04-04 to 2378-04-22 |
| 1 1 | 0 | 0x300000000 | `0x300000000` - `0x37fffffff` | 2378-04-22 to 2446-05-10 |

This is a somewhat odd encoding since there are effectively seven times as many positive values as negative values. There have also been long-standing bugs decoding and encoding dates beyond 2038, which don't seem to be fixed as of kernel 3.12 and e2fsprogs 1.42.8. 64-bit kernels incorrectly use the extra epoch bits 1,1 for dates between 1901 and 1970. At some point the kernel will be fixed and e2fsck will fix this situation, assuming that it is run before 2310.

### The Contents of inode.i_block

Depending on the type of file an inode describes, the 60 bytes of storage in `inode.i_block` can be used in different ways. In general, regular files and directories will use it for file block indexing information, and special files will use it for special purposes.

### Symbolic Links

The target of a symbolic link will be stored in this field if the target string is less than 60 bytes long. Otherwise, either extents or block maps will be used to allocate data blocks to store the link target.

### Direct/Indirect Block Addressing

In ext2/3, file block numbers were mapped to logical block numbers by means of an (up to) three level 1-1 block map. To find the logical block that stores a particular file block, the code would navigate through this increasingly complicated structure. Notice that there is neither a magic number nor a checksum to provide any level of confidence that the block isn't full of garbage.

[Table omitted because LaTeX doesn't support nested tables.]

Note that with this block mapping scheme, it is necessary to fill out a lot of mapping data even for a large contiguous file! This inefficiency led to the creation of the extent mapping scheme, discussed below.

Notice also that a file using this mapping scheme cannot be placed higher than 2^32 blocks.

### Extent Tree

In ext4, the file to logical block map has been replaced with an extent tree. Under the old scheme, allocating a contiguous run of 1,000 blocks requires an indirect block to map all 1,000 entries; with extents, the mapping is reduced to a single `struct ext4_extent` with `ee_len = 1000`. If flex_bg is enabled, it is possible to allocate very large files with a single extent, at a considerable reduction in metadata block use, and some improvement in disk efficiency. The inode must have the extents flag (0x80000) flag set for this feature to be in use.

Extents are arranged as a tree. Each node of the tree begins with a `struct ext4_extent_header`. If the node is an interior node (`eh.eh_depth` > 0), the header is followed by `eh.eh_entries` instances of `struct ext4_extent_idx`; each of these index entries points to a block containing more nodes in the extent tree. If the node is a leaf node (`eh.eh_depth == 0`), then the header is followed by `eh.eh_entries` instances of `struct ext4_extent`; these

instances point to the file's data blocks. The root node of the extent tree is stored in `inode.i_block`, which allows for the first four extents to be recorded without the use of extra metadata blocks.

The extent tree header is recorded in `struct ext4_extent_header`, which is 12 bytes long:

| Offset | Size | Name | Description |
|---|---|---|---|
| 0x0 | __le16 | eh_magic | Magic number, 0xF30A. |
| 0x2 | __le16 | eh_entries | Number of valid entries following the header. |
| 0x4 | __le16 | eh_max | Maximum number of entries that could follow the header. |
| 0x6 | __le16 | eh_depth | Depth of this extent node in the extent tree. 0 = this extent node points to data blocks; otherwise, this extent node points to other extent nodes. The extent tree can be at most 5 levels deep: a logical block number can be at most `2^32`, and the smallest `n` that satisfies `4*(((blocksize - 12)/12)^n) >= 2^32` is 5. |
| 0x8 | __le32 | eh_generation | Generation of the tree. (Used by Lustre, but not standard ext4). |

Internal nodes of the extent tree, also known as index nodes, are recorded as `struct ext4_extent_idx`, and are 12 bytes long:

| Offset | Size | Name | Description |
|---|---|---|---|
| 0x0 | __le32 | ei_block | This index node covers file blocks from 'block' onward. |
| 0x4 | __le32 | ei_leaf_lo | Lower 32-bits of the block number of the extent node that is the next level lower in the tree. The tree node pointed to can be either another internal node or a leaf node, described below. |
| 0x8 | __le16 | ei_leaf_hi | Upper 16-bits of the previous field. |
| 0xA | __u16 | ei_unused | |

Leaf nodes of the extent tree are recorded as `struct ext4_extent`, and are also 12 bytes long:

| Offset | Size | Name | Description |
|---|---|---|---|
| 0x0 | __le32 | ee_block | First file block number that this extent covers. |
| 0x4 | __le16 | ee_len | Number of blocks covered by extent. If the value of this field is <= 32768, the extent is initialized. If the value of the field is > 32768, the extent is uninitialized and the actual extent length is `ee_len` - 32768. Therefore, the maximum length of a initialized extent is 32768 blocks, and the maximum length of an uninitialized extent is 32767. |
| 0x6 | __le16 | ee_start_hi | Upper 16-bits of the block number to which this extent points. |
| 0x8 | __le32 | ee_start_lo | Lower 32-bits of the block number to which this extent points. |

Prior to the introduction of metadata checksums, the extent header + extent entries always left at least 4 bytes of unallocated space at the end of each extent tree data block (because (2^x % 12) >= 4). Therefore, the 32-bit checksum is inserted into this space. The 4 extents in the inode do not need checksumming, since the inode is already checksummed. The checksum is calculated against the FS UUID, the inode number, the inode generation, and the entire extent block leading up to (but not including) the checksum itself.

`struct ext4_extent_tail` is 4 bytes long:

| Offset | Size | Name | Description |
|---|---|---|---|
| 0x0 | __le32 | eb_checksum | Checksum of the extent block, crc32c(uuid+inum+igeneration+extentblock) |

### Inline Data

If the inline data feature is enabled for the filesystem and the flag is set for the inode, it is possible that the first 60 bytes of the file data are stored here.

### Directory Entries

In an ext4 filesystem, a directory is more or less a flat file that maps an arbitrary byte string (usually ASCII) to an inode number on the filesystem. There can be many directory entries across the filesystem that reference the same inode number--these are known as hard links, and that is why hard links cannot reference files on other filesystems. As such, directory entries are found by reading the data block(s) associated with a directory file for the particular directory entry that is desired.

### Linear (Classic) Directories

By default, each directory lists its entries in an "almost-linear" array. I write "almost" because it's not a linear array in the memory sense because directory entries are not split across filesystem blocks. Therefore, it is more accurate to say that a directory is a series of data blocks and that each block contains a linear array of directory entries. The end of each per-block array is signified by reaching the end of the block; the last entry in the block has a record length that takes it all the way to the end of the block. The end of the entire directory is of course signified by reaching the end of the file. Unused directory entries are signified by inode = 0. By default the filesystem uses `struct ext4_dir_entry_2` for directory entries unless the "filetype" feature flag is not set, in which case it uses `struct ext4_dir_entry`.

The original directory entry format is `struct ext4_dir_entry`, which is at most 263 bytes long, though on disk you'll need to reference `dirent.rec_len` to know for sure.

| Offset | Size | Name | Description |
|---|---|---|---|
| 0x0 | __le32 | inode | Number of the inode that this directory entry points to. |
| 0x4 | __le16 | rec_len | Length of this directory entry. Must be a multiple of 4. |
| 0x6 | __le16 | name_len | Length of the file name. |
| 0x8 | char | name[EXT4_NAME_LEN] | File name. |

Since file names cannot be longer than 255 bytes, the new directory entry format shortens the name_len field and uses the space for a file type flag, probably to avoid having to load every inode during directory tree traversal. This format is `ext4_dir_entry_2`, which is at most 263 bytes long, though on disk you'll need to reference `dirent.rec_len` to know for sure.

| Offset | Size | Name | Description |
|---|---|---|---|
| 0x0 | __le32 | inode | Number of the inode that this directory entry points to. |
| 0x4 | __le16 | rec_len | Length of this directory entry. |
| 0x6 | __u8 | name_len | Length of the file name. |
| 0x7 | __u8 | file_type | File type code, see *ftype* table below. |
| 0x8 | char | name[EXT4_NAME_LEN] | File name. |

The directory file type is one of the following values:

| Value | Description |
|---|---|
| 0x0 | Unknown. |
| 0x1 | Regular file. |
| 0x2 | Directory. |
| 0x3 | Character device file. |
| 0x4 | Block device file. |
| 0x5 | FIFO. |
| 0x6 | Socket. |
| 0x7 | Symbolic link. |

To support directories that are both encrypted and casefolded directories, we must also include hash information in the directory entry. We append `ext4_extended_dir_entry_2` to

ext4_dir_entry_2 except for the entries for dot and dotdot, which are kept the same. The structure follows immediately after name and is included in the size listed by rec_len If a directory entry uses this extension, it may be up to 271 bytes.

| Offset | Size | Name | Description |
|---|---|---|---|
| 0x0 | __le32 | hash | The hash of the directory name |
| 0x4 | __le32 | minor_hash | The minor hash of the directory name |

In order to add checksums to these classic directory blocks, a phony struct ext4_dir_entry is placed at the end of each leaf block to hold the checksum. The directory entry is 12 bytes long. The inode number and name_len fields are set to zero to fool old software into ignoring an apparently empty directory entry, and the checksum is stored in the place where the name normally goes. The structure is struct ext4_dir_entry_tail:

| Offset | Size | Name | Description |
|---|---|---|---|
| 0x0 | __le32 | det_reserved_zero1 | Inode number, which must be zero. |
| 0x4 | __le16 | det_rec_len | Length of this directory entry, which must be 12. |
| 0x6 | __u8 | det_reserved_zero2 | Length of the file name, which must be zero. |
| 0x7 | __u8 | det_reserved_ft | File type, which must be 0xDE. |
| 0x8 | __le32 | det_checksum | Directory leaf block checksum. |

The leaf directory block checksum is calculated against the FS UUID, the directory's inode number, the directory's inode generation number, and the entire directory entry block up to (but not including) the fake directory entry.

### Hash Tree Directories

A linear array of directory entries isn't great for performance, so a new feature was added to ext3 to provide a faster (but peculiar) balanced tree keyed off a hash of the directory entry name. If the EXT4_INDEX_FL (0x1000) flag is set in the inode, this directory uses a hashed btree (htree) to organize and find directory entries. For backwards read-only compatibility with ext2, this tree is actually hidden inside the directory file, masquerading as “empty” directory data blocks! It was stated previously that the end of the linear directory entry table was signified with an entry pointing to inode 0; this is (ab)used to fool the old linear-scan algorithm into thinking that the rest of the directory block is empty so that it moves on.

The root of the tree always lives in the first data block of the directory. By ext2 custom, the '.' and '..' entries must appear at the beginning of this first block, so they are put here as two struct ext4_dir_entry_2 s and not stored in the tree. The rest of the root node contains metadata about the tree and finally a hash->block map to find nodes that are lower in the htree. If dx_root.info.indirect_levels is non-zero then the htree has two levels; the data block pointed to by the root node's map is an interior node, which is indexed by a minor hash. Interior nodes in this tree contains a zeroed out struct ext4_dir_entry_2 followed by a minor_hash->block map to find leafe nodes. Leaf nodes contain a linear array of all struct ext4_dir_entry_2; all of these entries (presumably) hash to the same value. If there is an overflow, the entries simply overflow into the next leaf node, and the least-significant bit of the hash (in the interior node map) that gets us to this next leaf node is set.

To traverse the directory as a htree, the code calculates the hash of the desired file name and

uses it to find the corresponding block number. If the tree is flat, the block is a linear array of directory entries that can be searched; otherwise, the minor hash of the file name is computed and used against this second block to find the corresponding third block number. That third block number will be a linear array of directory entries.

To traverse the directory as a linear array (such as the old code does), the code simply reads every data block in the directory. The blocks used for the htree will appear to have no entries (aside from '.' and '..') and so only the leaf nodes will appear to have any interesting content.

The root of the htree is in `struct dx_root`, which is the full length of a data block:

| Offset | Type | Name | Description |
|---|---|---|---|
| 0x0 | __le32 | dot.inode | inode number of this directory. |
| 0x4 | __le16 | dot.rec_len | Length of this record, 12. |
| 0x6 | u8 | dot.name_len | Length of the name, 1. |
| 0x7 | u8 | dot.file_type | File type of this entry, 0x2 (directory) (if the feature flag is set). |
| 0x8 | char | dot.name[4] | ".000" |
| 0xC | __le32 | dotdot.inode | inode number of parent directory. |
| 0x10 | __le16 | dotdot.rec_len | block_size - 12. The record length is long enough to cover all htree data. |
| 0x12 | u8 | dotdot.name_len | Length of the name, 2. |
| 0x13 | u8 | dotdot.file_type | File type of this entry, 0x2 (directory) (if the feature flag is set). |
| 0x14 | char | dotdot_name[4] | "..00" |
| 0x18 | __le32 | struct dx_root_info.reserved_zero | Zero. |
| 0x1C | u8 | struct dx_root_info.hash_version | Hash type, see *dirhash* table below. |
| 0x1D | u8 | struct dx_root_info.info_length | Length of the tree information, 0x8. |
| 0x1E | u8 | struct dx_root_info.indirect_levels | Depth of the htree. Cannot be larger than 3 if the INCOMPAT_LARGEDIR feature is set; cannot be larger than 2 otherwise. |
| 0x1F | u8 | struct dx_root_info.unused_flags | |
| 0x20 | __le16 | limit | Maximum number of dx_entries that can follow this header, plus 1 for the header itself. |
| 0x22 | __le16 | count | Actual number of dx_entries that follow this header, plus 1 for the header itself. |
| 0x24 | __le32 | block | The block number (within the directory file) that goes with hash=0. |
| 0x28 | struct dx_entry | entries[0] | As many 8-byte `struct dx_entry` as fits in the rest of the data block. |

The directory hash is one of the following values:

| Value | Description |
|---|---|
| 0x0 | Legacy. |
| 0x1 | Half MD4. |
| 0x2 | Tea. |
| 0x3 | Legacy, unsigned. |
| 0x4 | Half MD4, unsigned. |
| 0x5 | Tea, unsigned. |
| 0x6 | Siphash. |

Interior nodes of an htree are recorded as `struct dx_node`, which is also the full length of a data block:

| Offset | Type | Name | Description |
|---|---|---|---|
| 0x0 | __le32 | fake.inode | Zero, to make it look like this entry is not in use. |
| 0x4 | __le16 | fake.rec_len | The size of the block, in order to hide all of the dx_node data. |
| 0x6 | u8 | name_len | Zero. There is no name for this “unused” directory entry. |
| 0x7 | u8 | file_type | Zero. There is no file type for this “unused” directory entry. |
| 0x8 | __le16 | limit | Maximum number of dx_entries that can follow this header, plus 1 for the header itself. |
| 0xA | __le16 | count | Actual number of dx_entries that follow this header, plus 1 for the header itself. |
| 0xE | __le32 | block | The block number (within the directory file) that goes with the lowest hash value of this block. This value is stored in the parent block. |
| 0x12 | struct dx_entry | entries[0] | As many 8-byte `struct dx_entry` as fits in the rest of the data block. |

The hash maps that exist in both `struct dx_root` and `struct dx_node` are recorded as `struct dx_entry`, which is 8 bytes long:

| Offset | Type | Name | Description |
|---|---|---|---|
| 0x0 | __le32 | hash | Hash code. |
| 0x4 | __le32 | block | Block number (within the directory file, not filesystem blocks) of the next node in the htree. |

(If you think this is all quite clever and peculiar, so does the author.)

If metadata checksums are enabled, the last 8 bytes of the directory block (precisely the length of one dx_entry) are used to store a `struct dx_tail`, which contains the checksum. The `limit` and `count` entries in the dx_root/dx_node structures are adjusted as necessary to fit the dx_tail into the block. If there is no space for the dx_tail, the user is notified to run e2fsck -D to rebuild the directory index (which will ensure that there's space for the checksum. The dx_tail structure is 8 bytes long and looks like this:

| Offset | Type | Name | Description |
|---|---|---|---|
| 0x0 | u32 | dt_reserved | Zero. |
| 0x4 | __le32 | dt_checksum | Checksum of the htree directory block. |

The checksum is calculated against the FS UUID, the htree index header (dx_root or dx_node), all of the htree indices (dx_entry) that are in use, and the tail block (dx_tail).

### Extended Attributes

Extended attributes (xattrs) are typically stored in a separate data block on the disk and referenced from inodes via `inode.i_file_acl*`. The first use of extended attributes seems to have been for storing file ACLs and other security data (selinux). With the `user_xattr` mount option it is possible for users to store extended attributes so long as all attribute names begin with "user"; this restriction seems to have disappeared as of Linux 3.0.

There are two places where extended attributes can be found. The first place is between the end of each inode entry and the beginning of the next inode entry. For example, if inode.i_extra_isize = 28 and sb.inode_size = 256, then there are 256 - (128 + 28) = 100 bytes available for in-inode extended attribute storage. The second place where extended attributes can be found is in the block pointed to by `inode.i_file_acl`. As of Linux 3.11, it is not possible for this block to contain a pointer to a second extended attribute block (or even the remaining blocks of a cluster). In theory it is possible for each attribute's value to be stored in a separate data block, though as of Linux 3.11 the code does not permit this.

Keys are generally assumed to be ASCIIZ strings, whereas values can be strings or binary data.

Extended attributes, when stored after the inode, have a header `ext4_xattr_ibody_header` that is 4 bytes long:

| Offset | Type | Name | Description |
|---|---|---|---|
| 0x0 | __le32 | h_magic | Magic number for identification, 0xEA020000. This value is set by the Linux driver, though e2fsprogs doesn't seem to check it(?) |

The beginning of an extended attribute block is in `struct ext4_xattr_header`, which is 32 bytes long:

| Offset | Type | Name | Description |
|---|---|---|---|
| 0x0 | __le32 | h_magic | Magic number for identification, 0xEA020000. |
| 0x4 | __le32 | h_refcount | Reference count. |
| 0x8 | __le32 | h_blocks | Number of disk blocks used. |
| 0xC | __le32 | h_hash | Hash value of all attributes. |
| 0x10 | __le32 | h_checksum | Checksum of the extended attribute block. |
| 0x14 | __u32 | h_reserved[3] | Zero. |

The checksum is calculated against the FS UUID, the 64-bit block number of the extended attribute block, and the entire block (header + entries).

Following the `struct ext4_xattr_header` or `struct ext4_xattr_ibody_header` is an array of `struct ext4_xattr_entry`; each of these entries is at least 16 bytes long. When stored in an external block, the `struct ext4_xattr_entry` entries must be stored in sorted order. The sort order is `e_name_index`, then `e_name_len`, and finally `e_name`. Attributes stored inside an inode do not need be stored in sorted order.

| Offset | Type | Name | Description |
|---|---|---|---|
| 0x0 | __u8 | e_name_len | Length of name. |
| 0x1 | __u8 | e_name_index | Attribute name index. There is a discussion of this below. |
| 0x2 | __le16 | e_value_offs | Location of this attribute's value on the disk block where it is stored. Multiple attributes can share the same value. For an inode attribute this value is relative to the start of the first entry; for a block this value is relative to the start of the block (i.e. the header). |
| 0x4 | __le32 | e_value_inum | The inode where the value is stored. Zero indicates the value is in the same block as this entry. This field is only used if the INCOMPAT_EA_INODE feature is enabled. |
| 0x8 | __le32 | e_value_size | Length of attribute value. |
| 0xC | __le32 | e_hash | Hash value of attribute name and attribute value. The kernel doesn't update the hash for in-inode attributes, so for that case this value must be zero, because e2fsck validates any non-zero hash regardless of where the xattr lives. |
| 0x10 | char | e_name[e_name_len] | Attribute name. Does not include trailing NULL. |

Attribute values can follow the end of the entry table. There appears to be a requirement that they be aligned to 4-byte boundaries. The values are stored starting at the end of the block and grow towards the xattr_header/xattr_entry table. When the two collide, the overflow is put into a separate disk block. If the disk block fills up, the filesystem returns -ENOSPC.

The first four fields of the `ext4_xattr_entry` are set to zero to mark the end of the key list.

### Attribute Name Indices

Logically speaking, extended attributes are a series of key=value pairs. The keys are assumed to be NULL-terminated strings. To reduce the amount of on-disk space that the keys consume, the beginning of the key string is matched against the attribute name index. If a match is found, the attribute name index field is set, and matching string is removed from the key name. Here is a map of name index values to key prefixes:

| Name Index | Key Prefix |
|---|---|
| 0 | (no prefix) |
| 1 | “user.” |
| 2 | “system.posix_acl_access” |
| 3 | “system.posix_acl_default” |
| 4 | “trusted.” |
| 6 | “security.” |
| 7 | “system.” (inline_data only?) |
| 8 | “system.richacl” (SuSE kernels only?) |

For example, if the attribute key is “user.fubar”, the attribute name index is set to 1 and the “fubar” name is recorded on disk.

### POSIX ACLs

POSIX ACLs are stored in a reduced version of the Linux kernel (and libacl's) internal ACL format. The key difference is that the version number is different (1) and the `e_id` field is only stored for named user and group ACLs.

## 3.23 WHAT IS Flash-Friendly File System (F2FS)?

NAND flash memory-based storage devices, such as SSD, eMMC, and SD cards, have been equipped on a variety systems ranging from mobile to server systems. Since they are known to have different characteristics from the conventional rotating disks, a file system, an upper layer to the storage device, should adapt to the changes from the sketch in the design level.

F2FS is a file system exploiting NAND flash memory-based storage devices, which is based on Log-structured File System (LFS). The design has been focused on addressing the fundamental issues in LFS, which are snowball effect of wandering tree and high cleaning overhead.

Since a NAND flash memory-based storage device shows different characteristic according to its internal geometry or flash memory management scheme, namely FTL, F2FS and its tools support various parameters not only for configuring on-disk layout, but also for selecting allocation and cleaning algorithms.

The following git tree provides the file system formatting tool (mkfs.f2fs), a consistency checking tool (fsck.f2fs), and a debugging tool (dump.f2fs).

- git://git.kernel.org/pub/scm/linux/kernel/git/jaegeuk/f2fs-tools.git

For sending patches, please use the following mailing list:

- linux-f2fs-devel@lists.sourceforge.net

For reporting bugs, please use the following f2fs bug tracker link:

- https://bugzilla.kernel.org/enter_bug.cgi?product=File%20System&component=f2fs

### 3.23.1 Background and Design issues

#### Log-structured File System (LFS)

”A log-structured file system writes all modifications to disk sequentially in a log-like structure, thereby speeding up both file writing and crash recovery. The log is the only structure on disk; it contains indexing information so that files can be read back from the log efficiently. In order to maintain large free areas on disk for fast writing, we divide the log into segments and use a segment cleaner to compress the live information from heavily fragmented segments.” from Rosenblum, M. and Ousterhout, J. K., 1992, ”The design and implementation of a log-structured file system”, ACM Trans. Computer Systems 10, 1, 26–52.

#### Wandering Tree Problem

In LFS, when a file data is updated and written to the end of log, its direct pointer block is updated due to the changed location. Then the indirect pointer block is also updated due to the direct pointer block update. In this manner, the upper index structures such as inode, inode map, and checkpoint block are also updated recursively. This problem is called as wandering tree problem [1], and in order to enhance the performance, it should eliminate or relax the update propagation as much as possible.

[1] Bityutskiy, A. 2005. JFFS3 design issues. http://www.linux-mtd.infradead.org/

#### Cleaning Overhead

Since LFS is based on out-of-place writes, it produces so many obsolete blocks scattered across the whole storage. In order to serve new empty log space, it needs to reclaim these obsolete blocks seamlessly to users. This job is called as a cleaning process.

The process consists of three operations as follows.

1. A victim segment is selected through referencing segment usage table.
2. It loads parent index structures of all the data in the victim identified by segment summary blocks.
3. It checks the cross-reference between the data and its parent index structure.
4. It moves valid data selectively.

This cleaning job may cause unexpected long delays, so the most important goal is to hide the latencies to users. And also definitely, it should reduce the amount of valid data to be moved, and move them quickly as well.

### 3.23.2 Key Features

#### Flash Awareness

- Enlarge the random write area for better performance, but provide the high spatial locality
- Align FS data structures to the operational units in FTL as best efforts

#### Wandering Tree Problem

- Use a term, “node”, that represents inodes as well as various pointer blocks
- Introduce Node Address Table (NAT) containing the locations of all the “node” blocks; this will cut off the update propagation.

#### Cleaning Overhead

- Support a background cleaning process
- Support greedy and cost-benefit algorithms for victim selection policies
- Support multi-head logs for static/dynamic hot and cold data separation
- Introduce adaptive logging for efficient block allocation

### 3.23.3 Mount Options

| | |
|---|---|
| background_gc=%s | Turn on/off cleaning operations, namely garbage collection, triggered in background when I/O subsystem is idle. If background_gc=on, it will turn on the garbage collection and if background_gc=off, garbage collection will be turned off. If background_gc=sync, it will turn on synchronous garbage collection running in background. Default value for this option is on. So garbage collection is on by default. |
| gc_merge | When background_gc is on, this option can be enabled to let background GC thread to handle foreground GC requests, it can eliminate the sluggish issue caused by slow foreground GC operation when GC is triggered from a process with limited I/O and CPU resources. |
| nogc_merge | Disable GC merge feature. |
| disable_roll_forward | Disable the roll-forward recovery routine |
| norecovery | Disable the roll-forward recovery routine, mounted read- only (i.e., -o ro,disable_roll_forward) |

continues on next page

Table 4 – continued from previous page

| | |
|---|---|
| discard/nodiscard | Enable/disable real-time discard in f2fs, if discard is enabled, f2fs will issue discard/TRIM commands when a segment is cleaned. |
| no_heap | Disable heap-style segment allocation which finds free segments for data from the beginning of main area, while for node from the end of main area. |
| nouser_xattr | Disable Extended User Attributes. Note: xattr is enabled by default if CONFIG_F2FS_FS_XATTR is selected. |
| noacl | Disable POSIX Access Control List. Note: acl is enabled by default if CONFIG_F2FS_FS_POSIX_ACL is selected. |
| active_logs=%u | Support configuring the number of active logs. In the current design, f2fs supports only 2, 4, and 6 logs. Default number is 6. |
| disable_ext_identify | Disable the extension list configured by mkfs, so f2fs is not aware of cold files such as media files. |
| inline_xattr | Enable the inline xattrs feature. |
| noinline_xattr | Disable the inline xattrs feature. |
| inline_xattr_size=%u | Support configuring inline xattr size, it depends on flexible inline xattr feature. |
| inline_data | Enable the inline data feature: Newly created small (<~3.4k) files can be written into inode block. |
| inline_dentry | Enable the inline dir feature: data in newly created directory entries can be written into inode block. The space of inode block which is used to store inline dentries is limited to ~3.4k. |
| noinline_dentry | Disable the inline dentry feature. |
| flush_merge | Merge concurrent cache_flush commands as much as possible to eliminate redundant command issues. If the underlying device handles the cache_flush command relatively slowly, recommend to enable this option. |
| nobarrier | This option can be used if underlying storage guarantees its cached data should be written to the novolatile area. If this option is set, no cache_flush commands are issued but f2fs still guarantees the write ordering of all the data writes. |
| barrier | If this option is set, cache_flush commands are allowed to be issued. |
| fastboot | This option is used when a system wants to reduce mount time as much as possible, even though normal performance can be sacrificed. |

continues on next page

Table 4 – continued from previous page

| | |
|---|---|
| extent_cache | Enable an extent cache based on rb-tree, it can cache as many as extent which map between contiguous logical address and physical address per inode, resulting in increasing the cache hit ratio. Set by default. |
| noextent_cache | Disable an extent cache based on rb-tree explicitly, see the above extent_cache mount option. |
| noinline_data | Disable the inline data feature, inline data feature is enabled by default. |
| data_flush | Enable data flushing before checkpoint in order to persist data of regular and symlink. |
| reserve_root=%d | Support configuring reserved space which is used for allocation from a privileged user with specified uid or gid, unit: 4KB, the default limit is 0.2% of user blocks. |
| resuid=%d | The user ID which may use the reserved blocks. |
| resgid=%d | The group ID which may use the reserved blocks. |
| fault_injection=%d | Enable fault injection in all supported types with specified injection rate. |

continues on next page

Table 4 – continued from previous page

<table>
<tr>
<td>fault_type=%d</td>
<td>Support configuring fault injection type, should be enabled with fault_injection option, fault type value is shown below, it supports single or combined type.
<table>
<tr><th>Type_Name</th><th>Type_Value</th></tr>
<tr><td>FAULT_KMALLOC</td><td>0x000000001</td></tr>
<tr><td>FAULT_KVMALLOC</td><td>0x000000002</td></tr>
<tr><td>FAULT_PAGE_ALLOC</td><td>0x000000004</td></tr>
<tr><td>FAULT_PAGE_GET</td><td>0x000000008</td></tr>
<tr><td>FAULT_ALLOC_BIO</td><td>0x000000010 (obsolete)</td></tr>
<tr><td>FAULT_ALLOC_NID</td><td>0x000000020</td></tr>
<tr><td>FAULT_ORPHAN</td><td>0x000000040</td></tr>
<tr><td>FAULT_BLOCK</td><td>0x000000080</td></tr>
<tr><td>FAULT_DIR_DEPTH</td><td>0x000000100</td></tr>
<tr><td>FAULT_EVICT_INODE</td><td>0x000000200</td></tr>
<tr><td>FAULT_TRUNCATE</td><td>0x000000400</td></tr>
<tr><td>FAULT_READ_IO</td><td>0x000000800</td></tr>
<tr><td>FAULT_CHECKPOINT</td><td>0x000001000</td></tr>
<tr><td>FAULT_DISCARD</td><td>0x000002000</td></tr>
<tr><td>FAULT_WRITE_IO</td><td>0x000004000</td></tr>
<tr><td>FAULT_SLAB_ALLOC</td><td>0x000008000</td></tr>
<tr><td>FAULT_DQUOT_INIT</td><td>0x000010000</td></tr>
<tr><td>FAULT_LOCK_OP</td><td>0x000020000</td></tr>
<tr><td>FAULT_BLKADDR</td><td>0x000040000</td></tr>
</table>
</td>
</tr>
</table>

continues on next page

Table 4 – continued from previous page

| | |
|---|---|
| mode=%s | Control block allocation mode which supports ”adaptive” and ”lfs”. In ”lfs” mode, there should be no random writes towards main area. ”fragment:segment” and ”fragment:block” are newly added here. These are developer options for experiments to simulate filesystem fragmentation/after-GC situation itself. The developers use these modes to understand filesystem fragmentation/after-GC condition well, and eventually get some insights to handle them better. In ”fragment:segment”, f2fs allocates a new segment in ramdom position. With this, we can simulate the after-GC condition. In ”fragment:block”, we can scatter block allocation with ”max_fragment_chunk” and ”max_fragment_hole” sysfs nodes. We added some randomness to both chunk and hole size to make it close to realistic IO pattern. So, in this mode, f2fs will allocate 1..<max_fragment_chunk> blocks in a chunk and make a hole in the length of 1..<max_fragment_hole> by turns. With this, the newly allocated blocks will be scattered throughout the whole partition. Note that ”fragment:block” implicitly enables ”fragment:segment” option for more randomness. Please, use these options for your experiments and we strongly recommend to reformat the filesystem after using these options. |
| io_bits=%u | Set the bit size of write IO requests. It should be set with ”mode=lfs”. |
| usrquota | Enable plain user disk quota accounting. |
| grpquota | Enable plain group disk quota accounting. |
| prjquota | Enable plain project quota accounting. |
| usrjquota=<file> | Appoint specified file and type during mount, so that quota |
| grpjquota=<file> | information can be properly updated during recovery flow, |
| prjjquota=<file> | <quota file>: must be in root directory; |
| jqfmt=<quota type> | <quota type>: [vfsold,vfsv0,vfsv1]. |
| offusrjquota | Turn off user journalled quota. |
| offgrpjquota | Turn off group journalled quota. |
| offprjjquota | Turn off project journalled quota. |
| quota | Enable plain user disk quota accounting. |
| noquota | Disable all plain disk quota option. |
| alloc_mode=%s | Adjust block allocation policy, which supports ”reuse” and ”default”. |

continues on next page

Table 4 – continued from previous page

| | |
|---|---|
| fsync_mode=%s | Control the policy of fsync. Currently supports ”posix”, ”strict”, and ”nobarrier”. In ”posix” mode, which is default, fsync will follow POSIX semantics and does a light operation to improve the filesystem performance. In ”strict” mode, fsync will be heavy and behaves in line with xfs, ext4 and btrfs, where xfstest generic/342 will pass, but the performance will regress. ”nobarrier” is based on ”posix”, but doesn't issue flush command for non-atomic files likewise ”nobarrier” mount option. |
| test_dummy_encryption | |
| test_dummy_encryption=%s | Enable dummy encryption, which provides a fake fscrypt context. The fake fscrypt context is used by xfstests. The argument may be either ”v1” or ”v2”, in order to select the corresponding fscrypt policy version. |
| checkpoint=%s[:%u[%]] | Set to ”disable” to turn off checkpointing. Set to ”enable” to reenable checkpointing. Is enabled by default. While disabled, any unmounting or unexpected shutdowns will cause the filesystem contents to appear as they did when the filesystem was mounted with that option. While mounting with checkpoint=disable, the filesystem must run garbage collection to ensure that all available space can be used. If this takes too much time, the mount may return EAGAIN. You may optionally add a value to indicate how much of the disk you would be willing to temporarily give up to avoid additional garbage collection. This can be given as a number of blocks, or as a percent. For instance, mounting with checkpoint=disable:100% would always succeed, but it may hide up to all remaining free space. The actual space that would be unusable can be viewed at /sys/fs/f2fs/<disk>/unusable This space is reclaimed once checkpoint=enable. |

continues on next page

Table 4 – continued from previous page

| | |
|---|---|
| checkpoint_merge | When checkpoint is enabled, this can be used to create a kernel daemon and make it to merge concurrent checkpoint requests as much as possible to eliminate redundant checkpoint issues. Plus, we can eliminate the sluggish issue caused by slow checkpoint operation when the checkpoint is done in a process context in a cgroup having low i/o budget and cpu shares. To make this do better, we set the default i/o priority of the kernel daemon to ”3”, to give one higher priority than other kernel threads. This is the same way to give a I/O priority to the jbd2 journaling thread of ext4 filesystem. |
| nocheckpoint_merge | Disable checkpoint merge feature. |
| compress_algorithm=%s | Control compress algorithm, currently f2fs supports ”lzo”, ”lz4”, ”zstd” and ”lzo-rle” algorithm. |
| compress_algorithm=%s:%d | Control compress algorithm and its compress level, now, only ”lz4” and ”zstd” support compress level config. algorithm level range lz4 3 - 16 zstd 1 - 22 |
| compress_log_size=%u | Support configuring compress cluster size. The size will be 4KB * (1 << %u). The default and minimum sizes are 16KB. |
| compress_extension=%s | Support adding specified extension, so that f2fs can enable compression on those corresponding files, e.g. if all files with '.ext' has high compression rate, we can set the '.ext' on compression extension list and enable compression on these file by default rather than to enable it via ioctl. For other files, we can still enable compression via ioctl. Note that, there is one reserved special extension '*', it can be set to enable compression for all files. |

continues on next page

Table 4 – continued from previous page

| | |
|---|---|
| nocompress_extension=%s | Support adding specified extension, so that f2fs can disable compression on those corresponding files, just contrary to compression extension. If you know exactly which files cannot be compressed, you can use this. The same extension name can't appear in both compress and nocompress extension at the same time. If the compress extension specifies all files, the types specified by the nocompress extension will be treated as special cases and will not be compressed. Don't allow use '*' to specifie all file in nocompress extension. After add nocompress_extension, the priority should be: dir_flag < comp_extention,nocompress_extension < comp_file_flag,no_comp_file_flag. See more in compression sections. |
| compress_chksum | Support verifying chksum of raw data in compressed cluster. |
| compress_mode=%s | Control file compression mode. This supports ”fs” and ”user” modes. In ”fs” mode (default), f2fs does automatic compression on the compression enabled files. In ”user” mode, f2fs disables the automaic compression and gives the user discretion of choosing the target file and the timing. The user can do manual compression/decompression on the compression enabled files using ioctls. |
| compress_cache | Support to use address space of a filesystem managed inode to cache compressed block, in order to improve cache hit ratio of random read. |
| inlinecrypt | When possible, encrypt/decrypt the contents of encrypted files using the blk-crypto framework rather than filesystem-layer encryption. This allows the use of inline encryption hardware. The on-disk format is unaffected. For more details, see Documentation/block/inline-encryption.rst. |
| atgc | Enable age-threshold garbage collection, it provides high effectiveness and efficiency on background GC. |

continues on next page

Table 4 – continued from previous page

| | |
|---|---|
| discard_unit=%s | Control discard unit, the argument can be ”block”, ”segment” and ”section”, issued discard command's offset/size will be aligned to the unit, by default, ”discard_unit=block” is set, so that small discard functionality is enabled. For blkzoned device, ”discard_unit=section” will be set by default, it is helpful for large sized SMR or ZNS devices to reduce memory cost by getting rid of fs metadata supports small discard. |
| memory=%s | Control memory mode. This supports ”normal” and ”low” modes. ”low” mode is introduced to support low memory devices. Because of the nature of low memory devices, in this mode, f2fs will try to save memory sometimes by sacrificing performance. ”normal” mode is the default mode and same as before. |
| age_extent_cache | Enable an age extent cache based on rb-tree. It records data block update frequency of the extent per inode, in order to provide better temperature hints for data block allocation. |
| errors=%s | Specify f2fs behavior on critical errors. This supports modes: ”panic”, ”continue” and ”remount-ro”, respectively, trigger panic immediately, continue without doing anything, and remount the partition in read-only mode. By default it uses ”continue” mode. ========== ========== ========== ======== mode continue remount-ro panic ========== ========== ========== ======== ========== access ops normal normal N/A syscall errors -EIO -EROFS N/A mount option rw ro N/A pending dir write keep keep N/A pending non-dir write drop keep N/A pending node write drop keep N/A pending meta write keep keep N/A ========== ========== ========== ======== |

### 3.23.4 Debugfs Entries

/sys/kernel/debug/f2fs/ contains information about all the partitions mounted as f2fs. Each file shows the whole f2fs information.

/sys/kernel/debug/f2fs/status includes:

- major file system information managed by f2fs currently
- average SIT information about whole segments
- current memory footprint consumed by f2fs.

### 3.23.5 Sysfs Entries

Information about mounted f2fs file systems can be found in /sys/fs/f2fs. Each mounted filesystem will have a directory in /sys/fs/f2fs based on its device name (i.e., /sys/fs/f2fs/sda). The files in each per-device directory are shown in table below.

Files in /sys/fs/f2fs/<devname> (see also Documentation/ABI/testing/sysfs-fs-f2fs)

### 3.23.6 Usage

1. Download userland tools and compile them.
2. Skip, if f2fs was compiled statically inside kernel. Otherwise, insert the f2fs.ko module:

```
# insmod f2fs.ko
```

3. Create a directory to use when mounting:

```
# mkdir /mnt/f2fs
```

4. Format the block device, and then mount as f2fs:

```
# mkfs.f2fs -l label /dev/block_device
# mount -t f2fs /dev/block_device /mnt/f2fs
```

#### mkfs.f2fs

The mkfs.f2fs is for the use of formatting a partition as the f2fs filesystem, which builds a basic on-disk layout.

The quick options consist of:

| | |
|---|---|
| `-l [label]` | Give a volume label, up to 512 unicode name. |
| `-a [0 or 1]` | Split start location of each area for heap-based allocation.<br>1 is set by default, which performs this. |
| `-o [int]` | Set overprovision ratio in percent over volume size.<br>5 is set by default. |
| `-s [int]` | Set the number of segments per section.<br>1 is set by default. |
| `-z [int]` | Set the number of sections per zone.<br>1 is set by default. |
| `-e [str]` | Set basic extension list. e.g. ”mp3,gif,mov” |
| `-t [0 or 1]` | Disable discard command or not.<br>1 is set by default, which conducts discard. |

Note: please refer to the manpage of mkfs.f2fs(8) to get full option list.

### fsck.f2fs

The fsck.f2fs is a tool to check the consistency of an f2fs-formatted partition, which examines whether the filesystem metadata and user-made data are cross-referenced correctly or not. Note that, initial version of the tool does not fix any inconsistency.

The quick options consist of:

```
-d debug level [default:0]
```

Note: please refer to the manpage of fsck.f2fs(8) to get full option list.

### dump.f2fs

The dump.f2fs shows the information of specific inode and dumps SSA and SIT to file. Each file is dump_ssa and dump_sit.

The dump.f2fs is used to debug on-disk data structures of the f2fs filesystem. It shows on-disk inode information recognized by a given inode number, and is able to dump all the SSA and SIT entries into predefined files, ./dump_ssa and ./dump_sit respectively.

The options consist of:

```
-d debug level [default:0]
-i inode no (hex)
-s [SIT dump segno from #1~#2 (decimal), for all 0~-1]
-a [SSA dump segno from #1~#2 (decimal), for all 0~-1]
```

Examples:

```
# dump.f2fs -i [ino] /dev/sdx
# dump.f2fs -s 0~-1 /dev/sdx (SIT dump)
# dump.f2fs -a 0~-1 /dev/sdx (SSA dump)
```

Note: please refer to the manpage of dump.f2fs(8) to get full option list.

#### sload.f2fs

The sload.f2fs gives a way to insert files and directories in the existing disk image. This tool is useful when building f2fs images given compiled files.

Note: please refer to the manpage of sload.f2fs(8) to get full option list.

#### resize.f2fs

The resize.f2fs lets a user resize the f2fs-formatted disk image, while preserving all the files and directories stored in the image.

Note: please refer to the manpage of resize.f2fs(8) to get full option list.

#### defrag.f2fs

The defrag.f2fs can be used to defragment scattered written data as well as filesystem metadata across the disk. This can improve the write speed by giving more free consecutive space.

Note: please refer to the manpage of defrag.f2fs(8) to get full option list.

#### f2fs_io

The f2fs_io is a simple tool to issue various filesystem APIs as well as f2fs-specific ones, which is very useful for QA tests.

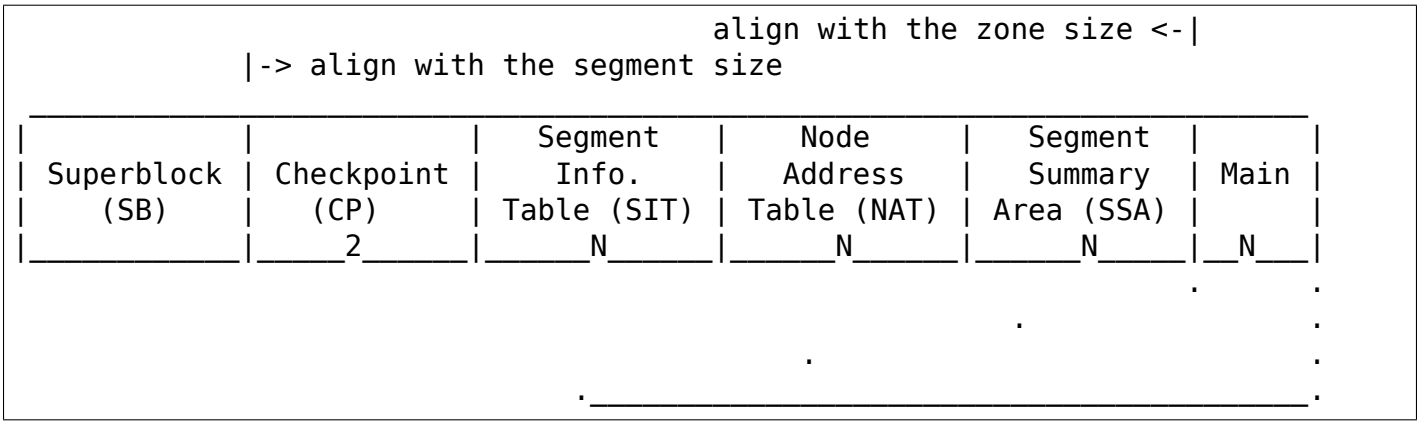
Note: please refer to the manpage of f2fs_io(8) to get full option list.

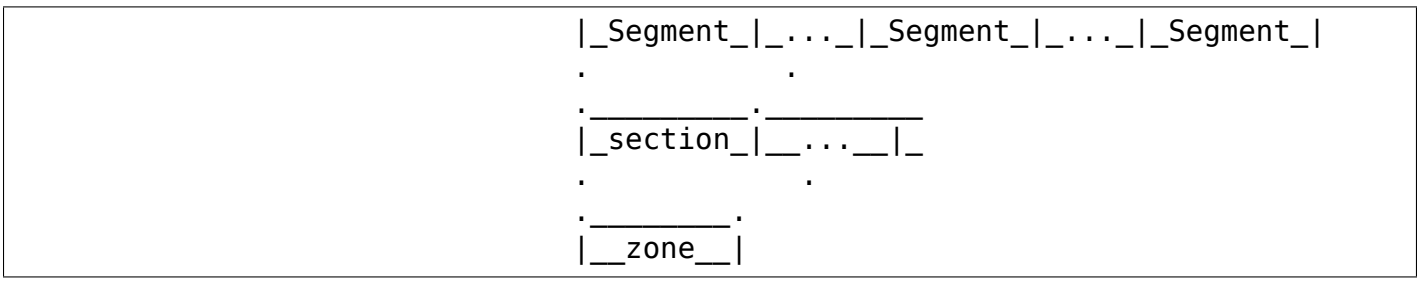
### 3.23.7 Design

#### On-disk Layout

F2FS divides the whole volume into a number of segments, each of which is fixed to 2MB in size. A section is composed of consecutive segments, and a zone consists of a set of sections. By default, section and zone sizes are set to one segment size identically, but users can easily modify the sizes by mkfs.

F2FS splits the entire volume into six areas, and all the areas except superblock consist of multiple segments as described below:






- **Superblock (SB)**
  It is located at the beginning of the partition, and there exist two copies to avoid file system crash. It contains basic partition information and some default parameters of f2fs.
- **Checkpoint (CP)**
  It contains file system information, bitmaps for valid NAT/SIT sets, orphan inode lists, and summary entries of current active segments.
- **Segment Information Table (SIT)**
  It contains segment information such as valid block count and bitmap for the validity of all the blocks.
- **Node Address Table (NAT)**
  It is composed of a block address table for all the node blocks stored in Main area.
- **Segment Summary Area (SSA)**
  It contains summary entries which contains the owner information of all the data and node blocks stored in Main area.
- **Main Area**
  It contains file and directory data including their indices.

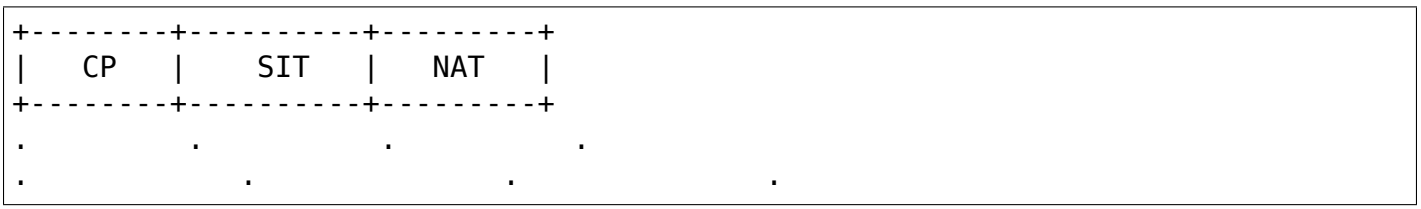
In order to avoid misalignment between file system and flash-based storage, F2FS aligns the start block address of CP with the segment size. Also, it aligns the start block address of Main area with the zone size by reserving some segments in SSA area.

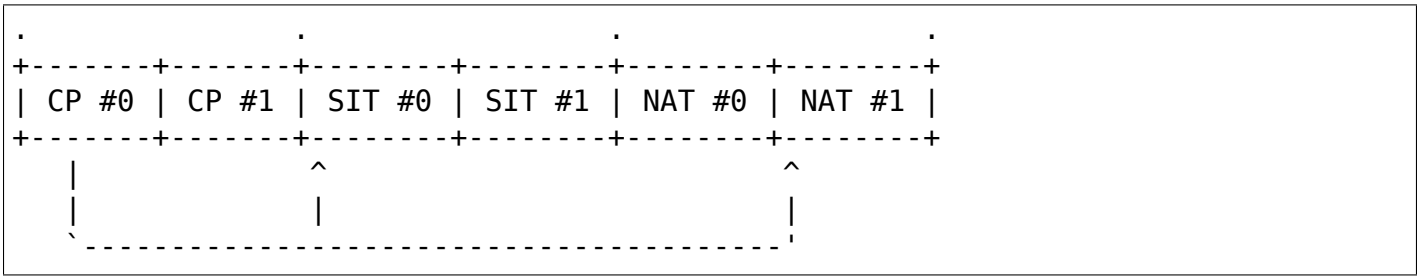
Reference the following survey for additional technical details. https://wiki.linaro.org/WorkingGroups/Kernel/Projects/FlashCardSurvey

### File System Metadata Structure

F2FS adopts the checkpointing scheme to maintain file system consistency. At mount time, F2FS first tries to find the last valid checkpoint data by scanning CP area. In order to reduce the scanning time, F2FS uses only two copies of CP. One of them always indicates the last valid data, which is called as shadow copy mechanism. In addition to CP, NAT and SIT also adopt the shadow copy mechanism.

For file system consistency, each CP points to which NAT and SIT copies are valid, as shown as below:






### Index Structure

The key data structure to manage the data locations is a ”node”. Similar to traditional file structures, F2FS has three types of node: inode, direct node, indirect node. F2FS assigns 4KB to an inode block which contains 923 data block indices, two direct node pointers, two indirect node pointers, and one double indirect node pointer as described below. One direct node block contains 1018 data blocks, and one indirect node block contains also 1018 node blocks. Thus, one inode block (i.e., a file) covers:

```
4KB * (923 + 2 * 1018 + 2 * 1018 * 1018 + 1018 * 1018 * 1018) := 3.94TB.

 Inode block (4KB)
   |- data (923)
   |- direct node (2)
   |          `- data (1018)
   |- indirect node (2)
   |            `- direct node (1018)
   |                        `- data (1018)
   `- double indirect node (1)
                       `- indirect node (1018)
                                   `- direct node (1018)
                                              `- data (1018)
```

Note that all the node blocks are mapped by NAT which means the location of each node is translated by the NAT table. In the consideration of the wandering tree problem, F2FS is able to cut off the propagation of node updates caused by leaf data writes.

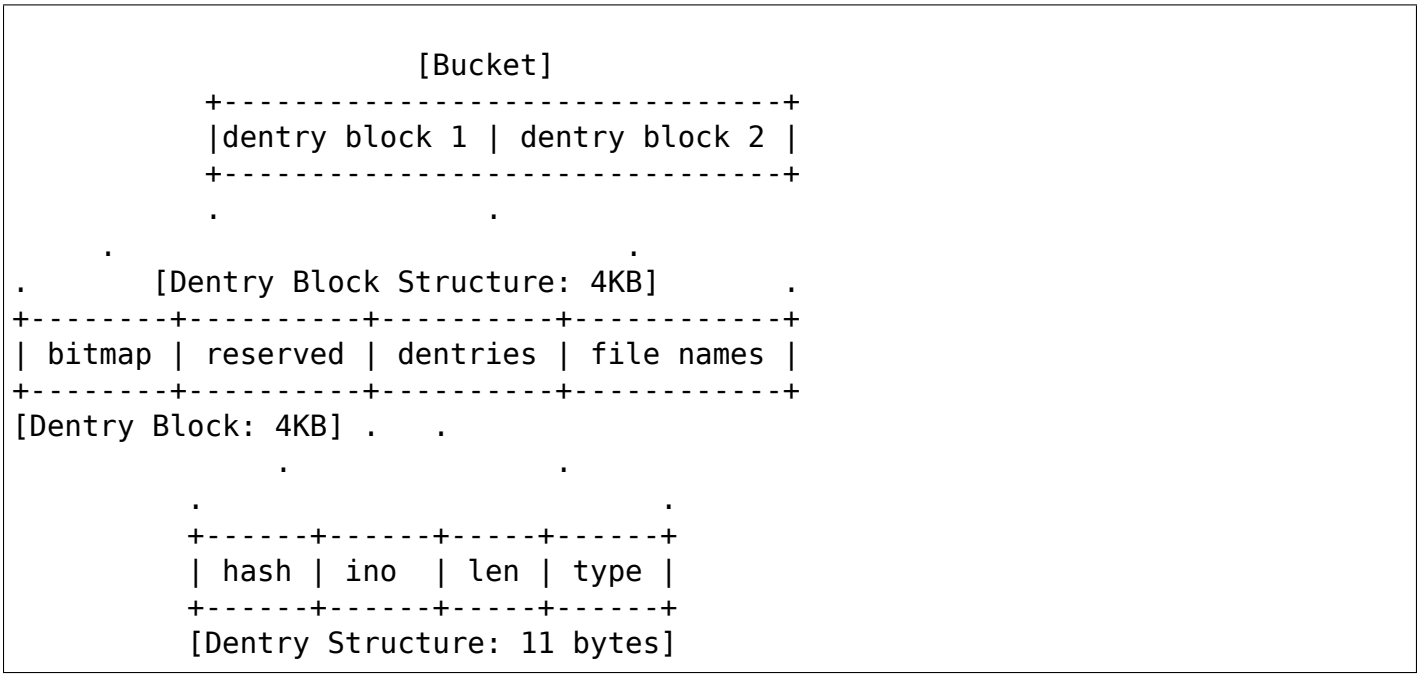
### Directory Structure

A directory entry occupies 11 bytes, which consists of the following attributes.

- hash hash value of the file name
- ino inode number
- len the length of file name
- type file type such as directory, symlink, etc

A dentry block consists of 214 dentry slots and file names. Therein a bitmap is used to represent whether each dentry is valid or not. A dentry block occupies 4KB with the following composition.

```
Dentry Block(4 K) = bitmap (27 bytes) + reserved (3 bytes) +
                    dentries(11 * 214 bytes) + file name (8 * 214 bytes)
```




F2FS implements multi-level hash tables for directory structure. Each level has a hash table with dedicated number of hash buckets as shown below. Note that "A(2B)" means a bucket includes 2 data blocks.

```
----------------------
A : bucket
B : block
N : MAX_DIR_HASH_DEPTH
----------------------

level #0   | A(2B)
        |
level #1   | A(2B) - A(2B)
        |
level #2   | A(2B) - A(2B) - A(2B) - A(2B)
    .      |   .       .       .       .
level #N/2 | A(2B) - A(2B) - A(2B) - A(2B) - A(2B) - ... - A(2B)
    .      |   .       .       .       .
level #N   | A(4B) - A(4B) - A(4B) - A(4B) - A(4B) - ... - A(4B)
```

The number of blocks and buckets are determined by:

```
                           ,- 2, if n < MAX_DIR_HASH_DEPTH / 2,
# of blocks in level #n = |
                           `- 4, Otherwise

                            ,- 2^(n + dir_level),
                            |        if n + dir_level < MAX_DIR_HASH_DEPTH / 2,
# of buckets in level #n = |
                            `- 2^((MAX_DIR_HASH_DEPTH / 2) - 1),
                                     Otherwise
```

When F2FS finds a file name in a directory, at first a hash value of the file name is calculated. Then, F2FS scans the hash table in level #0 to find the dentry consisting of the file name and its inode number. If not found, F2FS scans the next hash table in level #1. In this way, F2FS scans hash tables in each levels incrementally from 1 to N. In each level F2FS needs to scan only one bucket determined by the following equation, which shows O(log(# of files)) complexity:

```
bucket number to scan in level #n = (hash value) % (# of buckets in level #n)
```

In the case of file creation, F2FS finds empty consecutive slots that cover the file name. F2FS searches the empty slots in the hash tables of whole levels from 1 to N in the same way as the lookup operation.

The following figure shows an example of two cases holding children:

```
    --------------> Dir <--------------
    |                                  |
 child                               child

 child - child                       [hole] - child

 child - child - child               [hole] - [hole] - child

Case 1:                             Case 2:
Number of children = 6,             Number of children = 3,
File size = 7                       File size = 7
```

### Default Block Allocation

At runtime, F2FS manages six active logs inside ”Main” area: Hot/Warm/Cold node and Hot/Warm/Cold data.

- Hot node contains direct node blocks of directories.
- Warm node contains direct node blocks except hot node blocks.
- Cold node contains indirect node blocks
- Hot data contains dentry blocks
- Warm data contains data blocks except hot and cold data blocks
- Cold data contains multimedia data or migrated data blocks

LFS has two schemes for free space management: threaded log and copy-and-compac- tion. The copy-and-compaction scheme which is known as cleaning, is well-suited for devices showing very good sequential write performance, since free segments are served all the time for writing new data. However, it suffers from cleaning overhead under high utilization. Contrarily, the threaded log scheme suffers from random writes, but no cleaning process is needed. F2FS adopts a hybrid scheme where the copy-and-compaction scheme is adopted by default, but the policy is dynamically changed to the threaded log scheme according to the file system status.

In order to align F2FS with underlying flash-based storage, F2FS allocates a segment in a unit of section. F2FS expects that the section size would be the same as the unit size of garbage collection in FTL. Furthermore, with respect to the mapping granularity in FTL, F2FS allocates

each section of the active logs from different zones as much as possible, since FTL can write the data in the active logs into one allocation unit according to its mapping granularity.

### Cleaning process

F2FS does cleaning both on demand and in the background. On-demand cleaning is triggered when there are not enough free segments to serve VFS calls. Background cleaner is operated by a kernel thread, and triggers the cleaning job when the system is idle.

F2FS supports two victim selection policies: greedy and cost-benefit algorithms. In the greedy algorithm, F2FS selects a victim segment having the smallest number of valid blocks. In the cost-benefit algorithm, F2FS selects a victim segment according to the segment age and the number of valid blocks in order to address log block thrashing problem in the greedy algorithm. F2FS adopts the greedy algorithm for on-demand cleaner, while background cleaner adopts cost-benefit algorithm.

In order to identify whether the data in the victim segment are valid or not, F2FS manages a bitmap. Each bit represents the validity of a block, and the bitmap is composed of a bit stream covering whole blocks in main area.

### Fallocate(2) Policy

The default policy follows the below POSIX rule.

**Allocating disk space**
: The default operation (i.e., mode is zero) of fallocate() allocates the disk space within the range specified by offset and len. The file size (as reported by stat(2)) will be changed if offset+len is greater than the file size. Any subregion within the range specified by offset and len that did not contain data before the call will be initialized to zero. This default behavior closely resembles the behavior of the posix_fallocate(3) library function, and is intended as a method of optimally implementing that function.

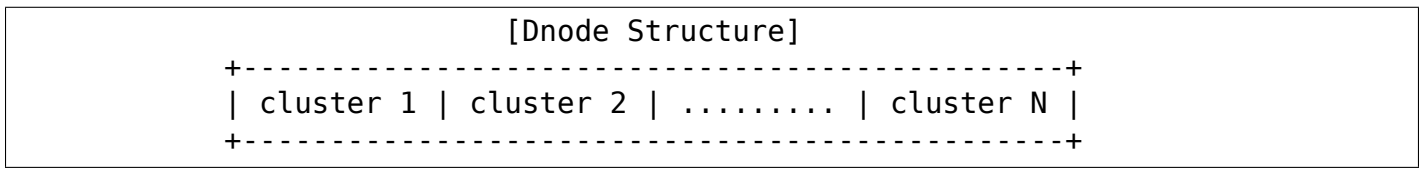
However, once F2FS receives ioctl(fd, F2FS_IOC_SET_PIN_FILE) in prior to fallocate(fd, DEFAULT_MODE), it allocates on-disk block addresses having zero or random data, which is useful to the below scenario where:

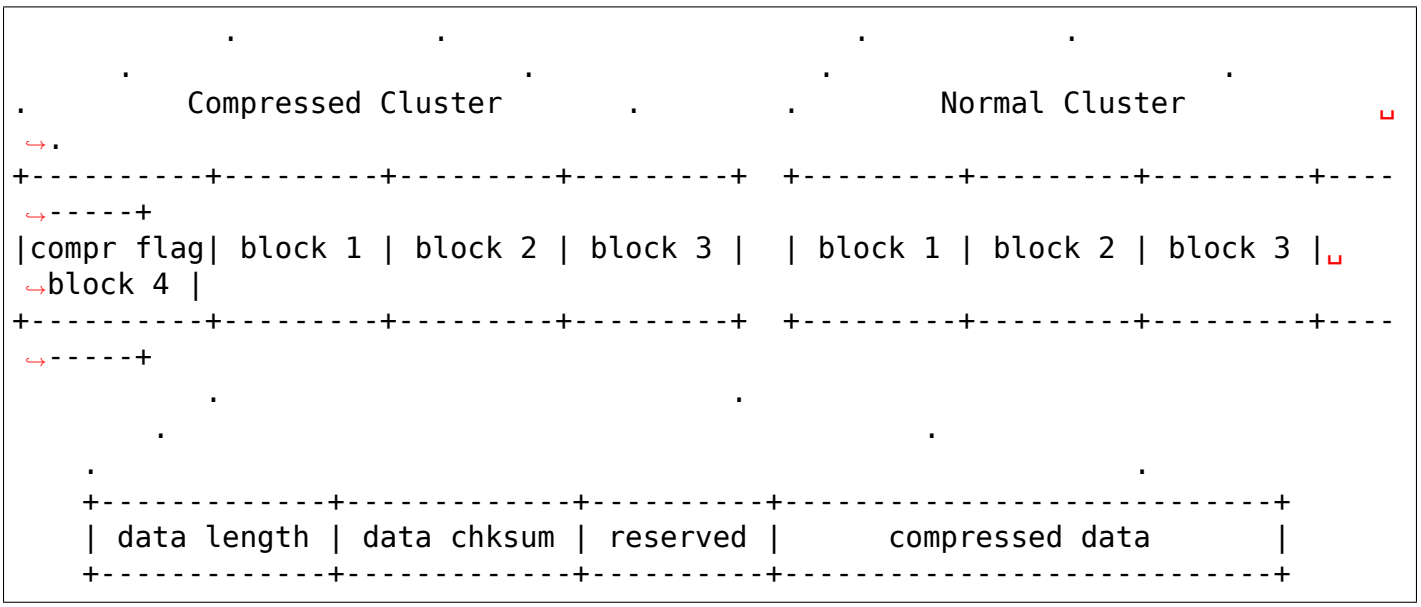
1. create(fd)
2. ioctl(fd, F2FS_IOC_SET_PIN_FILE)
3. fallocate(fd, 0, 0, size)
4. address = fibmap(fd, offset)
5. open(blkdev)
6. write(blkdev, address)

### Compression implementation

- New term named cluster is defined as basic unit of compression, file can be divided into multiple clusters logically. One cluster includes 4 << n (n >= 0) logical pages, compression size is also cluster size, each of cluster can be compressed or not.
- In cluster metadata layout, one special block address is used to indicate a cluster is a compressed one or normal one; for compressed cluster, following metadata maps cluster to [1, 4 << n - 1] physical blocks, in where f2fs stores data including compress header and compressed data.
- In order to eliminate write amplification during overwrite, F2FS only support compression on write-once file, data can be compressed only when all logical blocks in cluster contain valid data and compress ratio of cluster data is lower than specified threshold.
- To enable compression on regular inode, there are four ways:
  - chattr +c file
  - chattr +c dir; touch dir/file
  - mount w/ -o compress_extension=ext; touch file.ext
  - mount w/ -o compress_extension=*; touch any_file
- To disable compression on regular inode, there are two ways:
  - chattr -c file
  - mount w/ -o nocompress_extension=ext; touch file.ext
- Priority in between FS_COMPR_FL, FS_NOCOMP_FS, extensions:
  - compress_extension=so; nocompress_extension=zip; chattr +c dir; touch dir/foo.so; touch dir/bar.zip; touch dir/baz.txt; then foo.so and baz.txt should be compresse, bar.zip should be non-compressed. chattr +c dir/bar.zip can enable compress on bar.zip.
  - compress_extension=so; nocompress_extension=zip; chattr -c dir; touch dir/foo.so; touch dir/bar.zip; touch dir/baz.txt; then foo.so should be compresse, bar.zip and baz.txt should be non-compressed. chattr+c dir/bar.zip; chattr+c dir/baz.txt; can enable compress on bar.zip and baz.txt.
- At this point, compression feature doesn't expose compressed space to user directly in order to guarantee potential data updates later to the space. Instead, the main goal is to reduce data writes to flash disk as much as possible, resulting in extending disk life time as well as relaxing IO congestion. Alternatively, we've added ioctl(F2FS_IOC_RELEASE_COMPRESS_BLOCKS) interface to reclaim compressed space and show it to user after setting a special flag to the inode. Once the compressed space is released, the flag will block writing data to the file until either the compressed space is reserved via ioctl(F2FS_IOC_RESERVE_COMPRESS_BLOCKS) or the file size is truncated to zero.

Compress metadata layout:






### Compression mode

f2fs supports ”fs” and ”user” compression modes with ”compression_mode” mount option. With this option, f2fs provides a choice to select the way how to compress the compression enabled files (refer to ”Compression implementation” section for how to enable compression on a regular inode).

1) compress_mode=fs This is the default option. f2fs does automatic compression in the writeback of the compression enabled files.

2) compress_mode=user This disables the automatic compression and gives the user discretion of choosing the target file and the timing. The user can do manual compression/decompression on the compression enabled files using F2FS_IOC_DECOMPRESS_FILE and F2FS_IOC_COMPRESS_FILE ioctls like the below.

To decompress a file,

fd = open(filename, O_WRONLY, 0); ret = ioctl(fd, F2FS_IOC_DECOMPRESS_FILE);

To compress a file,

fd = open(filename, O_WRONLY, 0); ret = ioctl(fd, F2FS_IOC_COMPRESS_FILE);

### NVMe Zoned Namespace devices

- ZNS defines a per-zone capacity which can be equal or less than the zone-size. Zone-capacity is the number of usable blocks in the zone. F2FS checks if zone-capacity is less than zone-size, if it is, then any segment which starts after the zone-capacity is marked as not-free in the free segment bitmap at initial mount time. These segments are marked as permanently used so they are not allocated for writes and consequently are not needed to be garbage collected. In case the zone-capacity is not aligned to default segment size(2MB), then a segment can start before the zone-capacity and span across zone-capacity boundary. Such spanning segments are also considered as usable segments. All blocks past the zone-capacity are considered unusable in these segments.

## 3.24 Global File System 2

GFS2 is a cluster file system. It allows a cluster of computers to simultaneously use a block device that is shared between them (with FC, iSCSI, NBD, etc). GFS2 reads and writes to the block device like a local file system, but also uses a lock module to allow the computers coordinate their I/O so file system consistency is maintained. One of the nifty features of GFS2 is perfect consistency -- changes made to the file system on one machine show up immediately on all other machines in the cluster.

GFS2 uses interchangeable inter-node locking mechanisms, the currently supported mechanisms are:

**lock_nolock**

- allows GFS2 to be used as a local file system

**lock_dlm**

- uses the distributed lock manager (dlm) for inter-node locking. The dlm is found at linux/fs/dlm/

lock_dlm depends on user space cluster management systems found at the URL above.

To use GFS2 as a local file system, no external clustering systems are needed, simply:

```
$ mkfs -t gfs2 -p lock_nolock -j 1 /dev/block_device
$ mount -t gfs2 /dev/block_device /dir
```

The gfs2-utils package is required on all cluster nodes and, for lock_dlm, you will also need the dlm and corosync user space utilities configured as per the documentation.

gfs2-utils can be found at https://pagure.io/gfs2-utils

GFS2 is not on-disk compatible with previous versions of GFS, but it is pretty close.

The following man pages are available from gfs2-utils:

| | |
|---|---|
| fsck.gfs2 | to repair a filesystem |
| gfs2_grow | to expand a filesystem online |
| gfs2_jadd | to add journals to a filesystem online |
| tunegfs2 | to manipulate, examine and tune a filesystem |
| gfs2_convert | to convert a gfs filesystem to GFS2 in-place |
| mkfs.gfs2 | to make a filesystem |

## 3.25 uevents and GFS2

During the lifetime of a GFS2 mount, a number of uevents are generated. This document explains what the events are and what they are used for (by gfs_controld in gfs2-utils).

### 3.25.1 A list of GFS2 uevents

#### 1. ADD

The ADD event occurs at mount time. It will always be the first uevent generated by the newly created filesystem. If the mount is successful, an ONLINE uevent will follow. If it is not successful then a REMOVE uevent will follow.

The ADD uevent has two environment variables: SPECTATOR=[0|1] and RDONLY=[0|1] that specify the spectator status (a read-only mount with no journal assigned), and read-only (with journal assigned) status of the filesystem respectively.

#### 2. ONLINE

The ONLINE uevent is generated after a successful mount or remount. It has the same environment variables as the ADD uevent. The ONLINE uevent, along with the two environment variables for spectator and RDONLY are a relatively recent addition (2.6.32-rc+) and will not be generated by older kernels.

#### 3. CHANGE

The CHANGE uevent is used in two places. One is when reporting the successful mount of the filesystem by the first node (FIRSTMOUNT=Done). This is used as a signal by gfs_controld that it is then ok for other nodes in the cluster to mount the filesystem.

The other CHANGE uevent is used to inform of the completion of journal recovery for one of the filesystems journals. It has two environment variables, JID= which specifies the journal id which has just been recovered, and RECOVERY=[Done|Failed] to indicate the success (or otherwise) of the operation. These uevents are generated for every journal recovered, whether it is during the initial mount process or as the result of gfs_controld requesting a specific journal recovery via the /sys/fs/gfs2/<fsname>/lock_module/recovery file.

Because the CHANGE uevent was used (in early versions of gfs_controld) without checking the environment variables to discover the state, we cannot add any more functions to it without running the risk of someone using an older version of the user tools and breaking their cluster. For this reason the ONLINE uevent was used when adding a new uevent for a successful mount or remount.

#### 4. OFFLINE

The OFFLINE uevent is only generated due to filesystem errors and is used as part of the ”withdraw” mechanism. Currently this doesn't give any information about what the error is, which is something that needs to be fixed.

#### 5. REMOVE

The REMOVE uevent is generated at the end of an unsuccessful mount or at the end of a umount of the filesystem. All REMOVE uevents will have been preceded by at least an ADD uevent for the same filesystem, and unlike the other uevents is generated automatically by the kernel's kobject subsystem.

### 3.25.2 Information common to all GFS2 uevents (uevent environment variables)

#### 1. LOCKTABLE=

The LOCKTABLE is a string, as supplied on the mount command line (locktable=) or via fstab. It is used as a filesystem label as well as providing the information for a lock_dlm mount to be able to join the cluster.

#### 2. LOCKPROTO=

The LOCKPROTO is a string, and its value depends on what is set on the mount command line, or via fstab. It will be either lock_nolock or lock_dlm. In the future other lock managers may be supported.

#### 3. JOURNALID=

If a journal is in use by the filesystem (journals are not assigned for spectator mounts) then this will give the numeric journal id in all GFS2 uevents.

#### 4. UUID=

With recent versions of gfs2-utils, mkfs.gfs2 writes a UUID into the filesystem superblock. If it exists, this will be included in every uevent relating to the filesystem.

## 3.26 Glock internal locking rules

This documents the basic principles of the glock state machine internals. Each glock (struct gfs2_glock in fs/gfs2/incore.h) has two main (internal) locks:

1. A spinlock (gl_lockref.lock) which protects the internal state such as gl_state, gl_target and the list of holders (gl_holders)
2. A non-blocking bit lock, GLF_LOCK, which is used to prevent other threads from making calls to the DLM, etc. at the same time. If a thread takes this lock, it must then call run_queue (usually via the workqueue) when it releases it in order to ensure any pending tasks are completed.

The gl_holders list contains all the queued lock requests (not just the holders) associated with the glock. If there are any held locks, then they will be contiguous entries at the head of the list. Locks are granted in strictly the order that they are queued.

There are three lock states that users of the glock layer can request, namely shared (SH), deferred (DF) and exclusive (EX). Those translate to the following DLM lock modes:

| Glock mode | DLM | lock mode |
|---|---|---|
| UN | IV/NL | Unlocked (no DLM lock associated with glock) or NL |
| SH | PR | (Protected read) |
| DF | CW | (Concurrent write) |
| EX | EX | (Exclusive) |

Thus DF is basically a shared mode which is incompatible with the ”normal” shared lock mode, SH. In GFS2 the DF mode is used exclusively for direct I/O operations. The glocks are basically a lock plus some routines which deal with cache management. The following rules apply for the cache:

| Glock mode | Cache data | Cache Metadata | Dirty Data | Dirty Metadata |
|---|---|---|---|---|
| UN | No | No | No | No |
| SH | Yes | Yes | No | No |
| DF | No | Yes | No | No |
| EX | Yes | Yes | Yes | Yes |

These rules are implemented using the various glock operations which are defined for each type of glock. Not all types of glocks use all the modes. Only inode glocks use the DF mode for example.

Table of glock operations and per type constants:

| Field | Purpose |
|---|---|
| go_xmote_th | Called before remote state change (e.g. to sync dirty data) |
| go_xmote_bh | Called after remote state change (e.g. to refill cache) |
| go_inval | Called if remote state change requires invalidating the cache |
| go_demote_ok | Returns boolean value of whether its ok to demote a glock (e.g. checks time-out, and that there is no cached data) |
| go_lock | Called for the first local holder of a lock |
| go_unlock | Called on the final local unlock of a lock |
| go_dump | Called to print content of object for debugfs file, or on error to dump glock to the log. |
| go_type | The type of the glock, `LM_TYPE_*` |
| go_callback | Called if the DLM sends a callback to drop this lock |
| go_flags | GLOF_ASPACE is set, if the glock has an address space associated with it |

The minimum hold time for each lock is the time after a remote lock grant for which we ignore remote demote requests. This is in order to prevent a situation where locks are being bounced around the cluster from node to node with none of the nodes making any progress. This tends to show up most with shared mmapped files which are being written to by multiple nodes. By delaying the demotion in response to a remote callback, that gives the userspace program time to make some progress before the pages are unmapped.

There is a plan to try and remove the go_lock and go_unlock callbacks if possible, in order to try and speed up the fast path though the locking. Also, eventually we hope to make the glock ”EX” mode locally shared such that any local locking will be done with the i_mutex as required rather than via the glock.

Locking rules for glock operations:

| Operation | GLF_LOCK bit lock held | gl_lockref.lock spinlock held |
|---|---|---|
| go_xmote_th | Yes | No |
| go_xmote_bh | Yes | No |
| go_inval | Yes | No |
| go_demote_ok | Sometimes | Yes |
| go_lock | Yes | No |
| go_unlock | Yes | No |
| go_dump | Sometimes | Yes |
| go_callback | Sometimes (N/A) | Yes |

**Note:** Operations must not drop either the bit lock or the spinlock if its held on entry. go_dump and do_demote_ok must never block. Note that go_dump will only be called if the glock's state indicates that it is caching uptodate data.

Glock locking order within GFS2:

1. i_rwsem (if required)
2. Rename glock (for rename only)
3. Inode glock(s) (Parents before children, inodes at ”same level” with same parent in lock number order)
4. Rgrp glock(s) (for (de)allocation operations)
5. Transaction glock (via gfs2_trans_begin) for non-read operations
6. i_rw_mutex (if required)
7. Page lock (always last, very important!)

There are two glocks per inode. One deals with access to the inode itself (locking order as above), and the other, known as the iopen glock is used in conjunction with the i_nlink field in the inode to determine the lifetime of the inode in question. Locking of inodes is on a per-inode basis. Locking of rgrps is on a per rgrp basis. In general we prefer to lock local locks prior to cluster locks.

### 3.26.1 Glock Statistics

The stats are divided into two sets: those relating to the super block and those relating to an individual glock. The super block stats are done on a per cpu basis in order to try and reduce the overhead of gathering them. They are also further divided by glock type. All timings are in nanoseconds.

In the case of both the super block and glock statistics, the same information is gathered in each case. The super block timing statistics are used to provide default values for the glock timing statistics, so that newly created glocks should have, as far as possible, a sensible starting point. The per-glock counters are initialised to zero when the glock is created. The per-glock statistics are lost when the glock is ejected from memory.

The statistics are divided into three pairs of mean and variance, plus two counters. The mean/variance pairs are smoothed exponential estimates and the algorithm used is one which

will be very familiar to those used to calculation of round trip times in network code. See "TCP/IP Illustrated, Volume 1", W. Richard Stevens, sect 21.3, "Round-Trip Time Measurement", p. 299 and onwards. Also, Volume 2, Sect. 25.10, p. 838 and onwards. Unlike the TCP/IP Illustrated case, the mean and variance are not scaled, but are in units of integer nanoseconds.

The three pairs of mean/variance measure the following things:

1. DLM lock time (non-blocking requests)
2. DLM lock time (blocking requests)
3. Inter-request time (again to the DLM)

A non-blocking request is one which will complete right away, whatever the state of the DLM lock in question. That currently means any requests when (a) the current state of the lock is exclusive, i.e. a lock demotion (b) the requested state is either null or unlocked (again, a demotion) or (c) the "try lock" flag is set. A blocking request covers all the other lock requests.

There are two counters. The first is there primarily to show how many lock requests have been made, and thus how much data has gone into the mean/variance calculations. The other counter is counting queuing of holders at the top layer of the glock code. Hopefully that number will be a lot larger than the number of dlm lock requests issued.

So why gather these statistics? There are several reasons we'd like to get a better idea of these timings:

1. To be able to better set the glock "min hold time"
2. To spot performance issues more easily
3. To improve the algorithm for selecting resource groups for allocation (to base it on lock wait time, rather than blindly using a "try lock")

Due to the smoothing action of the updates, a step change in some input quantity being sampled will only fully be taken into account after 8 samples (or 4 for the variance) and this needs to be carefully considered when interpreting the results.

Knowing both the time it takes a lock request to complete and the average time between lock requests for a glock means we can compute the total percentage of the time for which the node is able to use a glock vs. time that the rest of the cluster has its share. That will be very useful when setting the lock min hold time.

Great care has been taken to ensure that we measure exactly the quantities that we want, as accurately as possible. There are always inaccuracies in any measuring system, but I hope this is as accurate as we can reasonably make it.

Per sb stats can be found here:

```
/sys/kernel/debug/gfs2/<fsname>/sbstats
```

Per glock stats can be found here:

```
/sys/kernel/debug/gfs2/<fsname>/glstats
```

Assuming that debugfs is mounted on /sys/kernel/debug and also that <fsname> is replaced with the name of the gfs2 filesystem in question.

The abbreviations used in the output as are follows:

| | |
|---|---|
| srtt | Smoothed round trip time for non blocking dlm requests |
| srttvar | Variance estimate for srtt |
| srttb | Smoothed round trip time for (potentially) blocking dlm requests |
| srttvarb | Variance estimate for srttb |
| sirt | Smoothed inter request time (for dlm requests) |
| sirtvar | Variance estimate for sirt |
| dlm | Number of dlm requests made (dcnt in glstats file) |
| queue | Number of glock requests queued (qcnt in glstats file) |

The sbstats file contains a set of these stats for each glock type (so 8 lines for each type) and for each cpu (one column per cpu). The glstats file contains a set of these stats for each glock in a similar format to the glocks file, but using the format mean/variance for each of the timing stats.

The gfs2_glock_lock_time tracepoint prints out the current values of the stats for the glock in question, along with some addition information on each dlm reply that is received:

| | |
|---|---|
| status | The status of the dlm request |
| flags | The dlm request flags |
| tdiff | The time taken by this specific request |

(remaining fields as per above list)

## 3.27 Macintosh HFS Filesystem for Linux

**Note:** This filesystem doesn't have a maintainer.

HFS stands for `Hierarchical File System` and is the filesystem used by the Mac Plus and all later Macintosh models. Earlier Macintosh models used MFS (`Macintosh File System`), which is not supported, MacOS 8.1 and newer support a filesystem called HFS+ that's similar to HFS but is extended in various areas. Use the hfsplus filesystem driver to access such filesystems from Linux.

### 3.27.1 Mount options

When mounting an HFS filesystem, the following options are accepted:

**creator=cccc, type=cccc**
Specifies the creator/type values as shown by the MacOS finder used for creating new files. Default values: '????'.

**uid=n, gid=n**
Specifies the user/group that owns all files on the filesystems. Default: user/group id of the mounting process.

**dir_umask=n, file_umask=n, umask=n**
Specifies the umask used for all files , all directories or all files and directories. Defaults to the umask of the mounting process.

**session=n**
Select the CDROM session to mount as HFS filesystem. Defaults to leaving that decision to the CDROM driver. This option will fail with anything but a CDROM as underlying devices.

**part=n**
Select partition number n from the devices. Does only makes sense for CDROMS because they can't be partitioned under Linux. For disk devices the generic partition parsing code does this for us. Defaults to not parsing the partition table at all.

**quiet**
Ignore invalid mount options instead of complaining.

### 3.27.2 Writing to HFS Filesystems

HFS is not a UNIX filesystem, thus it does not have the usual features you'd expect:

- You can't modify the set-uid, set-gid, sticky or executable bits or the uid and gid of files.
- You can't create hard- or symlinks, device files, sockets or FIFOs.

HFS does on the other have the concepts of multiple forks per file. These non-standard forks are represented as hidden additional files in the normal filesystems namespace which is kind of a cludge and makes the semantics for the a little strange:

- You can't create, delete or rename resource forks of files or the Finder's metadata.
- They are however created (with default values), deleted and renamed along with the corresponding data fork or directory.
- Copying files to a different filesystem will loose those attributes that are essential for MacOS to work.

### 3.27.3 Creating HFS filesystems

The hfsutils package from Robert Leslie contains a program called hformat that can be used to create HFS filesystem. See <https://www.mars.org/home/rob/proj/hfs/> for details.

### 3.27.4 Credits

The HFS drivers was written by Paul H. Hargrovea (hargrove@sccm.Stanford.EDU). Roman Zippel (roman@ardistech.com) rewrote large parts of the code and brought in btree routines derived from Brad Boyer's hfsplus driver.

## 3.28 Macintosh HFSPlus Filesystem for Linux

HFSPlus is a filesystem first introduced in MacOS 8.1. HFSPlus has several extensions to HFS, including 32-bit allocation blocks, 255-character unicode filenames, and file sizes of 2^63 bytes.

### 3.28.1 Mount options

When mounting an HFSPlus filesystem, the following options are accepted:

**creator=cccc, type=cccc**
Specifies the creator/type values as shown by the MacOS finder used for creating new files. Default values: '????'.

**uid=n, gid=n**
Specifies the user/group that owns all files on the filesystem that have uninitialized permissions structures. Default: user/group id of the mounting process.

**umask=n**
Specifies the umask (in octal) used for files and directories that have uninitialized permissions structures. Default: umask of the mounting process.

**session=n**
Select the CDROM session to mount as HFSPlus filesystem. Defaults to leaving that decision to the CDROM driver. This option will fail with anything but a CDROM as underlying devices.

**part=n**
Select partition number n from the devices. This option only makes sense for CDROMs because they can't be partitioned under Linux. For disk devices the generic partition parsing code does this for us. Defaults to not parsing the partition table at all.

**decompose**
Decompose file name characters.

**nodecompose**
Do not decompose file name characters.

**force**
Used to force write access to volumes that are marked as journalled or locked. Use at your own risk.

**nls=cccc**
Encoding to use when presenting file names.

### 3.28.2 References

kernel source: <file:fs/hfsplus>

Apple Technote 1150 https://developer.apple.com/legacy/library/technotes/tn/tn1150.html

## 3.29 Read/Write HPFS 2.09

1998-2004, Mikulas Patocka

**email**
mikulas@artax.karlin.mff.cuni.cz

**homepage**
https://artax.karlin.mff.cuni.cz/~mikulas/vyplody/hpfs/index-e.cgi

### 3.29.1 Credits

**Chris Smith, 1993, original read-only HPFS, some code and hpfs structures file**
is taken from it

Jacques Gelinas, MSDos mmap, Inspired by fs/nfs/mmap.c (Jon Tombs 15 Aug 1993)

Werner Almesberger, 1992, 1993, MSDos option parser & CR/LF conversion

Mount options

**uid=xxx,gid=xxx,umask=xxx (default uid=gid=0 umask=default_system_umask)**
Set owner/group/mode for files that do not have it specified in extended attributes. Mode is inverted umask - for example umask 027 gives owner all permission, group read permission and anybody else no access. Note that for files mode is anded with 0666. If you want files to have 'x' rights, you must use extended attributes.

**case=lower,asis (default asis)**
File name lowercasing in readdir.

**conv=binary,text,auto (default binary)**
CR/LF -> LF conversion, if auto, decision is made according to extension - there is a list of text extensions (I thing it's better to not convert text file than to damage binary file). If you want to change that list, change it in the source. Original readonly HPFS contained some strange heuristic algorithm that I removed. I thing it's danger to let the computer decide whether file is text or binary. For example, DJGPP binaries contain small text message at the beginning and they could be misidentified and damaged under some circumstances.

**check=none,normal,strict (default normal)**
Check level. Selecting none will cause only little speedup and big danger. I tried to write it so that it won't crash if check=normal on corrupted filesystems. check=strict means many superfluous checks - used for debugging (for example it checks if file is allocated in bitmaps when accessing it).

**errors=continue,remount-ro,panic (default remount-ro)**
Behaviour when filesystem errors found.

**chkdsk=no,errors,always (default errors)**
When to mark filesystem dirty so that OS/2 checks it.

**eas=no,ro,rw (default rw)**
: What to do with extended attributes. 'no' - ignore them and use always values specified in uid/gid/mode options. 'ro' - read extended attributes but do not create them. 'rw' - create extended attributes when you use chmod/chown/chgrp/mknod/ln -s on the filesystem.

**timeshift=(-)nnn (default 0)**
: Shifts the time by nnn seconds. For example, if you see under linux one hour more, than under os/2, use timeshift=-3600.

### 3.29.2 File names

As in OS/2, filenames are case insensitive. However, shell thinks that names are case sensitive, so for example when you create a file FOO, you can use 'cat FOO', 'cat Foo', 'cat foo' or 'cat F*' but not 'cat f*'. Note, that you also won't be able to compile linux kernel (and maybe other things) on HPFS because kernel creates different files with names like bootsect.S and bootsect.s. When searching for file thats name has characters >= 128, codepages are used - see below. OS/2 ignores dots and spaces at the end of file name, so this driver does as well. If you create 'a. ...', the file 'a' will be created, but you can still access it under names 'a.', 'a..', 'a . . . ' etc.

### 3.29.3 Extended attributes

On HPFS partitions, OS/2 can associate to each file a special information called extended attributes. Extended attributes are pairs of (key,value) where key is an ascii string identifying that attribute and value is any string of bytes of variable length. OS/2 stores window and icon positions and file types there. So why not use it for unix-specific info like file owner or access rights? This driver can do it. If you chown/chgrp/chmod on a hpfs partition, extended attributes with keys ”UID”, ”GID” or ”MODE” and 2-byte values are created. Only that extended attributes those value differs from defaults specified in mount options are created. Once created, the extended attributes are never deleted, they're just changed. It means that when your default uid=0 and you type something like 'chown luser file; chown root file' the file will contain extended attribute UID=0. And when you umount the fs and mount it again with uid=luser_uid, the file will be still owned by root! If you chmod file to 444, extended attribute ”MODE” will not be set, this special case is done by setting read-only flag. When you mknod a block or char device, besides ”MODE”, the special 4-byte extended attribute ”DEV” will be created containing the device number. Currently this driver cannot resize extended attributes - it means that if somebody (I don't know who?) has set ”UID”, ”GID”, ”MODE” or ”DEV” attributes with different sizes, they won't be rewritten and changing these values doesn't work.

### 3.29.4 Symlinks

You can do symlinks on HPFS partition, symlinks are achieved by setting extended attribute named ”SYMLINK” with symlink value. Like on ext2, you can chown and chgrp symlinks but I don't know what is it good for. chmoding symlink results in chmoding file where symlink points. These symlinks are just for Linux use and incompatible with OS/2. OS/2 PmShell symlinks are not supported because they are stored in very crazy way. They tried to do it so that link changes when file is moved ... sometimes it works. But the link is partly stored in directory extended attributes and partly in OS2SYS.INI. I don't want (and don't know how) to analyze or change OS2SYS.INI.

### 3.29.5 Codepages

HPFS can contain several uppercasing tables for several codepages and each file has a pointer to codepage its name is in. However OS/2 was created in America where people don't care much about codepages and so multiple codepages support is quite buggy. I have Czech OS/2 working in codepage 852 on my disk. Once I booted English OS/2 working in cp 850 and I created a file on my 852 partition. It marked file name codepage as 850 - good. But when I again booted Czech OS/2, the file was completely inaccessible under any name. It seems that OS/2 uppercases the search pattern with its system code page (852) and file name it's comparing to with its code page (850). These could never match. Is it really what IBM developers wanted? But problems continued. When I created in Czech OS/2 another file in that directory, that file was inaccessible too. OS/2 probably uses different uppercasing method when searching where to place a file (note, that files in HPFS directory must be sorted) and when searching for a file. Finally when I opened this directory in PmShell, PmShell crashed (the funny thing was that, when rebooted, PmShell tried to reopen this directory again :-). chkdsk happily ignores these errors and only low-level disk modification saved me. Never mix different language versions of OS/2 on one system although HPFS was designed to allow that. OK, I could implement complex codepage support to this driver but I think it would cause more problems than benefit with such buggy implementation in OS/2. So this driver simply uses first codepage it finds for uppercasing and lowercasing no matter what's file codepage index. Usually all file names are in this codepage - if you don't try to do what I described above :-)

### 3.29.6 Known bugs

HPFS386 on OS/2 server is not supported. HPFS386 installed on normal OS/2 client should work. If you have OS/2 server, use only read-only mode. I don't know how to handle some HPFS386 structures like access control list or extended perm list, I don't know how to delete them when file is deleted and how to not overwrite them with extended attributes. Send me some info on these structures and I'll make it. However, this driver should detect presence of HPFS386 structures, remount read-only and not destroy them (I hope).

When there's not enough space for extended attributes, they will be truncated and no error is returned.

OS/2 can't access files if the path is longer than about 256 chars but this driver allows you to do it. chkdsk ignores such errors.

Sometimes you won't be able to delete some files on a very full filesystem (returning error ENOSPC). That's because file in non-leaf node in directory tree (one directory, if it's large, has dirents in tree on HPFS) must be replaced with another node when deleted. And that new file might have larger name than the old one so the new name doesn't fit in directory node (dnode). And that would result in directory tree splitting, that takes disk space. Workaround is to delete other files that are leaf (probability that the file is non-leaf is about 1/50) or to truncate file first to make some space. You encounter this problem only if you have many directories so that preallocated directory band is full i.e.:

```
number_of_directories / size_of_filesystem_in_mb > 4.
```

You can't delete open directories.

You can't rename over directories (what is it good for?).

Renaming files so that only case changes doesn't work. This driver supports it but vfs doesn't. Something like 'mv file FILE' won't work.

All atimes and directory mtimes are not updated. That's because of performance reasons. If you extremely wish to update them, let me know, I'll write it (but it will be slow).

When the system is out of memory and swap, it may slightly corrupt filesystem (lost files, unbalanced directories). (I guess all filesystem may do it).

When compiled, you get warning: function declaration isn't a prototype. Does anybody know what does it mean?

### 3.29.7 What does ”unbalanced tree” message mean?

Old versions of this driver created sometimes unbalanced dnode trees. OS/2 chkdsk doesn't scream if the tree is unbalanced (and sometimes creates unbalanced trees too :-) but both HPFS and HPFS386 contain bug that it rarely crashes when the tree is not balanced. This driver handles unbalanced trees correctly and writes warning if it finds them. If you see this message, this is probably because of directories created with old version of this driver. Workaround is to move all files from that directory to another and then back again. Do it in Linux, not OS/2! If you see this message in directory that is whole created by this driver, it is BUG - let me know about it.

### 3.29.8 Bugs in OS/2

When you have two (or more) lost directories pointing each to other, chkdsk locks up when repairing filesystem.

Sometimes (I think it's random) when you create a file with one-char name under OS/2, OS/2 marks it as 'long'. chkdsk then removes this flag saying ”Minor fs error corrected”.

File names like ”a .b” are marked as 'long' by OS/2 but chkdsk ”corrects” it and marks them as short (and writes ”minor fs error corrected”). This bug is not in HPFS386.

### 3.29.9 Codepage bugs described above

If you don't install fixpacks, there are many, many more...

### 3.29.10 History

| | |
|---|---|
| 0.90 | First public release |
| 0.91 | Fixed bug that caused shooting to memory when write_inode was called on open inode (rarely happened) |
| 0.92 | Fixed a little memory leak in freeing directory inodes |
| 0.93 | Fixed bug that locked up the machine when there were too many filenames with first 15 characters same Fixed write_file to zero file when writing behind file end |
| 0.94 | Fixed a little memory leak when trying to delete busy file or directory |
| 0.95 | Fixed a bug that i_hpfs_parent_dir was not updated when moving files |
| 1.90 | First version for 2.1.1xx kernels |
| 1.91 | Fixed a bug that chk_sectors failed when sectors were at the end of disk Fixed a race-condition when write_inode is called while deleting file Fixed a bug that could possibly happen (with very low probability) when using 0xff in filenames.<br>Rewritten locking to avoid race-conditions<br>Mount option 'eas' now works<br>Fsync no longer returns error<br>Files beginning with '.' are marked hidden<br>Remount support added<br>Alloc is not so slow when filesystem becomes full<br>Atimes are no more updated because it slows down operation<br>Code cleanup (removed all commented debug prints) |
| 1.92 | Corrected a bug when sync was called just before closing file |
| 1.93 | Modified, so that it works with kernels >= 2.1.131, I don't know if it works with previous versions<br>Fixed a possible problem with disks > 64G (but I don't have one, so I can't test it)<br>Fixed a file overflow at 2G<br>Added new option 'timeshift'<br>Changed behaviour on HPFS386: It is now possible to operate on HPFS386 in read-only mode<br>Fixed a bug that slowed down alloc and prevented allocating 100% space (this bug was not destructive) |
| 1.94 | Added workaround for one bug in Linux<br>Fixed one buffer leak<br>Fixed some incompatibilities with large extended attributes (but it's still not 100% ok, I have no info on it and OS/2 doesn't want to create them)<br>Rewritten allocation<br>Fixed a bug with i_blocks (du sometimes didn't display correct values)<br>Directories have no longer archive attribute set (some programs don't like it)<br>Fixed a bug that it set badly one flag in large anode tree (it was not destructive) |
| 1.95 | Fixed one buffer leak, that could happen on corrupted filesystem<br>Fixed one bug in allocation in 1.94 |
| 1.96 | Added workaround for one bug in OS/2 (HPFS locked up, HPFS386 reported error sometimes when opening directories in PMSHELL)<br>Fixed a possible bitmap race<br>Fixed possible problem on large disks<br>You can now delete open files<br>Fixed a nondestructive race in rename |
| 1.97 | Support for HPFS v3 (on large partitions)<br>ZFixed a bug that it didn't allow creation of files > 128M (it should be 2G) |
| 1.97.1 | Changed names of global symbols<br>Fixed a bug when chmoding or chowning root directory |
| 1.98 | Fixed a deadlock when using old_readdir Better directory handling, workaround for ”unbalanced tree” bug in OS/2 |
| 1.99 | Corrected a possible problem when there's not enough space while deleting file<br>Now it tries to truncate the file if there's not enough space when deleting |

## 3.30 FUSE

### 3.30.1 Definitions

**Userspace filesystem:**
A filesystem in which data and metadata are provided by an ordinary userspace process. The filesystem can be accessed normally through the kernel interface.

**Filesystem daemon:**
The process(es) providing the data and metadata of the filesystem.

**Non-privileged mount (or user mount):**
A userspace filesystem mounted by a non-privileged (non-root) user. The filesystem daemon is running with the privileges of the mounting user. NOTE: this is not the same as mounts allowed with the "user" option in /etc/fstab, which is not discussed here.

**Filesystem connection:**
A connection between the filesystem daemon and the kernel. The connection exists until either the daemon dies, or the filesystem is umounted. Note that detaching (or lazy umounting) the filesystem does *not* break the connection, in this case it will exist until the last reference to the filesystem is released.

**Mount owner:**
The user who does the mounting.

**User:**
The user who is performing filesystem operations.

### 3.30.2 What is FUSE?

FUSE is a userspace filesystem framework. It consists of a kernel module (fuse.ko), a userspace library (libfuse.*) and a mount utility (fusermount).

One of the most important features of FUSE is allowing secure, non-privileged mounts. This opens up new possibilities for the use of filesystems. A good example is sshfs: a secure network filesystem using the sftp protocol.

The userspace library and utilities are available from the FUSE homepage:

### 3.30.3 Filesystem type

The filesystem type given to mount(2) can be one of the following:

**fuse**
This is the usual way to mount a FUSE filesystem. The first argument of the mount system call may contain an arbitrary string, which is not interpreted by the kernel.

**fuseblk**
The filesystem is block device based. The first argument of the mount system call is interpreted as the name of the device.

### 3.30.4 Mount options

**fd=N**
The file descriptor to use for communication between the userspace filesystem and the kernel. The file descriptor must have been obtained by opening the FUSE device ('/dev/fuse').

**rootmode=M**
The file mode of the filesystem's root in octal representation.

**user_id=N**
The numeric user id of the mount owner.

**group_id=N**
The numeric group id of the mount owner.

**default_permissions**
By default FUSE doesn't check file access permissions, the filesystem is free to implement its access policy or leave it to the underlying file access mechanism (e.g. in case of network filesystems). This option enables permission checking, restricting access based on file mode. It is usually useful together with the 'allow_other' mount option.

**allow_other**
This option overrides the security measure restricting file access to the user mounting the filesystem. This option is by default only allowed to root, but this restriction can be removed with a (userspace) configuration option.

**max_read=N**
With this option the maximum size of read operations can be set. The default is infinite. Note that the size of read requests is limited anyway to 32 pages (which is 128kbyte on i386).

**blksize=N**
Set the block size for the filesystem. The default is 512. This option is only valid for 'fuseblk' type mounts.

### 3.30.5 Control filesystem

There's a control filesystem for FUSE, which can be mounted by:

```
mount -t fusectl none /sys/fs/fuse/connections
```

Mounting it under the '/sys/fs/fuse/connections' directory makes it backwards compatible with earlier versions.

Under the fuse control filesystem each connection has a directory named by a unique number.

For each connection the following files exist within this directory:

**waiting**
The number of requests which are waiting to be transferred to userspace or being processed by the filesystem daemon. If there is no filesystem activity and 'waiting' is non-zero, then the filesystem is hung or deadlocked.

**abort**
Writing anything into this file will abort the filesystem connection. This means

that all waiting requests will be aborted an error returned for all aborted and new requests.

Only the owner of the mount may read or write these files.

#### Interrupting filesystem operations

If a process issuing a FUSE filesystem request is interrupted, the following will happen:

- If the request is not yet sent to userspace AND the signal is fatal (SIGKILL or unhandled fatal signal), then the request is dequeued and returns immediately.
- If the request is not yet sent to userspace AND the signal is not fatal, then an interrupted flag is set for the request. When the request has been successfully transferred to userspace and this flag is set, an INTERRUPT request is queued.
- If the request is already sent to userspace, then an INTERRUPT request is queued.

INTERRUPT requests take precedence over other requests, so the userspace filesystem will receive queued INTERRUPTs before any others.

The userspace filesystem may ignore the INTERRUPT requests entirely, or may honor them by sending a reply to the *original* request, with the error set to EINTR.

It is also possible that there's a race between processing the original request and its INTERRUPT request. There are two possibilities:

1. The INTERRUPT request is processed before the original request is processed
2. The INTERRUPT request is processed after the original request has been answered

If the filesystem cannot find the original request, it should wait for some timeout and/or a number of new requests to arrive, after which it should reply to the INTERRUPT request with an EAGAIN error. In case 1) the INTERRUPT request will be requeued. In case 2) the INTERRUPT reply will be ignored.

### 3.30.6 Aborting a filesystem connection

It is possible to get into certain situations where the filesystem is not responding. Reasons for this may be:

a) Broken userspace filesystem implementation

b) Network connection down

c) Accidental deadlock

d) Malicious deadlock

(For more on c) and d) see later sections)

In either of these cases it may be useful to abort the connection to the filesystem. There are several ways to do this:

- Kill the filesystem daemon. Works in case of a) and b)
- Kill the filesystem daemon and all users of the filesystem. Works in all cases except some malicious deadlocks

- Use forced umount (umount -f). Works in all cases but only if filesystem is still attached (it hasn't been lazy unmounted)
- Abort filesystem through the FUSE control filesystem. Most powerful method, always works.

### 3.30.7 How do non-privileged mounts work?

Since the mount() system call is a privileged operation, a helper program (fusermount) is needed, which is installed setuid root.

The implication of providing non-privileged mounts is that the mount owner must not be able to use this capability to compromise the system. Obvious requirements arising from this are:

A) mount owner should not be able to get elevated privileges with the help of the mounted filesystem

B) mount owner should not get illegitimate access to information from other users' and the super user's processes

C) mount owner should not be able to induce undesired behavior in other users' or the super user's processes

### 3.30.8 How are requirements fulfilled?

A) The mount owner could gain elevated privileges by either:

   1. creating a filesystem containing a device file, then opening this device
   2. creating a filesystem containing a suid or sgid application, then executing this application

   The solution is not to allow opening device files and ignore setuid and setgid bits when executing programs. To ensure this fusermount always adds ”nosuid” and ”nodev” to the mount options for non-privileged mounts.

B) If another user is accessing files or directories in the filesystem, the filesystem daemon serving requests can record the exact sequence and timing of operations performed. This information is otherwise inaccessible to the mount owner, so this counts as an information leak.

   The solution to this problem will be presented in point 2) of C).

C) There are several ways in which the mount owner can induce undesired behavior in other users' processes, such as:

   1) mounting a filesystem over a file or directory which the mount owner could otherwise not be able to modify (or could only make limited modifications).

      This is solved in fusermount, by checking the access permissions on the mountpoint and only allowing the mount if the mount owner can do unlimited modification (has write access to the mountpoint, and mountpoint is not a ”sticky” directory)

   2) Even if 1) is solved the mount owner can change the behavior of other users' processes.

      i) It can slow down or indefinitely delay the execution of a filesystem operation creating a DoS against the user or the whole system. For example a suid application

locking a system file, and then accessing a file on the mount owner's filesystem could be stopped, and thus causing the system file to be locked forever.

ii) It can present files or directories of unlimited length, or directory structures of unlimited depth, possibly causing a system process to eat up diskspace, memory or other resources, again causing *DoS*.

The solution to this as well as B) is not to allow processes to access the filesystem, which could otherwise not be monitored or manipulated by the mount owner. Since if the mount owner can ptrace a process, it can do all of the above without using a FUSE mount, the same criteria as used in ptrace can be used to check if a process is allowed to access the filesystem or not.

Note that the *ptrace* check is not strictly necessary to prevent C/2/i, it is enough to check if mount owner has enough privilege to send signal to the process accessing the filesystem, since *SIGSTOP* can be used to get a similar effect.

### 3.30.9 I think these limitations are unacceptable?

If a sysadmin trusts the users enough, or can ensure through other measures, that system processes will never enter non-privileged mounts, it can relax the last limitation in several ways:

- With the 'user_allow_other' config option. If this config option is set, the mounting user can add the 'allow_other' mount option which disables the check for other users' processes.

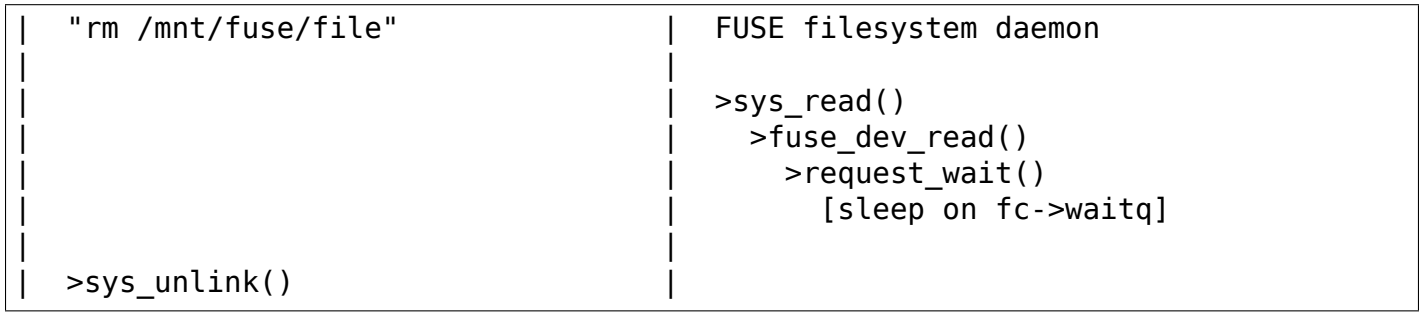
  User namespaces have an unintuitive interaction with 'allow_other': an unprivileged user - normally restricted from mounting with 'allow_other' - could do so in a user namespace where they're privileged. If any process could access such an 'allow_other' mount this would give the mounting user the ability to manipulate processes in user namespaces where they're unprivileged. For this reason 'allow_other' restricts access to users in the same userns or a descendant.

- With the 'allow_sys_admin_access' module option. If this option is set, super user's processes have unrestricted access to mounts irrespective of allow_other setting or user namespace of the mounting user.

Note that both of these relaxations expose the system to potential information leak or *DoS* as described in points B and C/2/i-ii in the preceding section.

### 3.30.10 Kernel - userspace interface

The following diagram shows how a filesystem operation (in this example unlink) is performed in FUSE.



```
|    >fuse_unlink()                   |                                    |
|      [get request from              |                                    |
|       fc->unused_list]              |                                    |
|      >request_send()                |                                    |
|        [queue req on fc->pending]   |                                    |
|        [wake up fc->waitq]          |        [woken up]                  |
|        >request_wait_answer()       |                                    |
|          [sleep on req->waitq]      |                                    |
|                                     |      <request_wait()               |
|                                     |      [remove req from fc->pending] |
|                                     |      [copy req to read buffer]     |
|                                     |      [add req to fc->processing]   |
|                                     |    <fuse_dev_read()                |
|                                     |  <sys_read()                       |
|                                     |                                    |
|                                     |  [perform unlink]                  |
|                                     |                                    |
|                                     |  >sys_write()                      |
|                                     |    >fuse_dev_write()               |
|                                     |      [look up req in fc->processing]|
|                                     |      [remove from fc->processing]  |
|                                     |      [copy write buffer to req]    |
|          [woken up]                 |      [wake up req->waitq]          |
|                                     |    <fuse_dev_write()               |
|                                     |  <sys_write()                      |
|        <request_wait_answer()       |                                    |
|      <request_send()                |                                    |
|      [add request to                |                                    |
|       fc->unused_list]              |                                    |
|    <fuse_unlink()                   |                                    |
|  <sys_unlink()                      |                                    |
```

**Note:** Everything in the description above is greatly simplified

There are a couple of ways in which to deadlock a FUSE filesystem. Since we are talking about unprivileged userspace programs, something must be done about these.

**Scenario 1 - Simple deadlock**:

```
|  "rm /mnt/fuse/file"                |  FUSE filesystem daemon            |
|                                     |                                    |
|  >sys_unlink("/mnt/fuse/file")      |                                    |
|    [acquire inode semaphore         |                                    |
|     for "file"]                     |                                    |
|    >fuse_unlink()                   |                                    |
|      [sleep on req->waitq]          |                                    |
|                                     |  <sys_read()                       |
|                                     |  >sys_unlink("/mnt/fuse/file")     |
|                                     |    [acquire inode semaphore        |
|                                     |     for "file"]                    |
```

```
|                                       |    *DEADLOCK*
```

The solution for this is to allow the filesystem to be aborted.

**Scenario 2 - Tricky deadlock**

This one needs a carefully crafted filesystem. It's a variation on the above, only the call back to the filesystem is not explicit, but is caused by a pagefault.

```
|  Kamikaze filesystem thread 1       |  Kamikaze filesystem thread 2
|                                     |
|  [fd = open("/mnt/fuse/file")]      |  [request served normally]
|  [mmap fd to 'addr']                |
|  [close fd]                         |  [FLUSH triggers 'magic' flag]
|  [read a byte from addr]            |
|    >do_page_fault()                 |
|      [find or create page]          |
|      [lock page]                    |
|      >fuse_readpage()               |
|         [queue READ request]        |
|         [sleep on req->waitq]       |
|                                     |  [read request to buffer]
|                                     |  [create reply header before addr]
|                                     |  >sys_write(addr - headerlength)
|                                     |    >fuse_dev_write()
|                                     |      [look up req in fc->processing]
|                                     |      [remove from fc->processing]
|                                     |      [copy write buffer to req]
|                                     |        >do_page_fault()
|                                     |           [find or create page]
|                                     |           [lock page]
|                                     |           * DEADLOCK *
```

The solution is basically the same as above.

An additional problem is that while the write buffer is being copied to the request, the request must not be interrupted/aborted. This is because the destination address of the copy may not be valid after the request has returned.

This is solved with doing the copy atomically, and allowing abort while the page(s) belonging to the write buffer are faulted with get_user_pages(). The 'req->locked' flag indicates when the copy is taking place, and abort is delayed until this flag is unset.

## 3.31 Fuse I/O Modes

Fuse supports the following I/O modes:

- direct-io
- cached + write-through + writeback-cache

The direct-io mode can be selected with the FOPEN_DIRECT_IO flag in the FUSE_OPEN reply.

In direct-io mode the page cache is completely bypassed for reads and writes. No read-ahead takes place. Shared mmap is disabled by default. To allow shared mmap, the FUSE_DIRECT_IO_ALLOW_MMAP flag may be enabled in the FUSE_INIT reply.

In cached mode reads may be satisfied from the page cache, and data may be read-ahead by the kernel to fill the cache. The cache is always kept consistent after any writes to the file. All mmap modes are supported.

The cached mode has two sub modes controlling how writes are handled. The write-through mode is the default and is supported on all kernels. The writeback-cache mode may be selected by the FUSE_WRITEBACK_CACHE flag in the FUSE_INIT reply.

In write-through mode each write is immediately sent to userspace as one or more WRITE requests, as well as updating any cached pages (and caching previously uncached, but fully written pages). No READ requests are ever sent for writes, so when an uncached page is partially written, the page is discarded.

In writeback-cache mode (enabled by the FUSE_WRITEBACK_CACHE flag) writes go to the cache only, which means that the write(2) syscall can often complete very fast. Dirty pages are written back implicitly (background writeback or page reclaim on memory pressure) or explicitly (invoked by close(2), fsync(2) and when the last ref to the file is being released on munmap(2)). This mode assumes that all changes to the filesystem go through the FUSE kernel module (size and atime/ctime/mtime attributes are kept up-to-date by the kernel), so it's generally not suitable for network filesystems. If a partial page is written, then the page needs to be first read from userspace. This means, that even for files opened for O_WRONLY it is possible that READ requests will be generated by the kernel.

## 3.32 Inotify - A Powerful yet Simple File Change Notification System

Document started 15 Mar 2005 by Robert Love <rml@novell.com>

Document updated 4 Jan 2015 by Zhang Zhen <zhenzhang.zhang@huawei.com>

- Deleted obsoleted interface, just refer to manpages for user interface.

(i) Rationale

**Q:**

What is the design decision behind not tying the watch to the open fd of the watched object?

**A:**

Watches are associated with an open inotify device, not an open file. This solves the primary problem with dnotify: keeping the file open pins the file and thus, worse, pins the

mount. Dnotify is therefore infeasible for use on a desktop system with removable media as the media cannot be unmounted. Watching a file should not require that it be open.

**Q:**

What is the design decision behind using an-fd-per-instance as opposed to an fd-per-watch?

**A:**

An fd-per-watch quickly consumes more file descriptors than are allowed, more fd's than are feasible to manage, and more fd's than are optimally select()-able. Yes, root can bump the per-process fd limit and yes, users can use epoll, but requiring both is a silly and extraneous requirement. A watch consumes less memory than an open file, separating the number spaces is thus sensible. The current design is what user-space developers want: Users initialize inotify, once, and add n watches, requiring but one fd and no twiddling with fd limits. Initializing an inotify instance two thousand times is silly. If we can implement user-space's preferences cleanly--and we can, the idr layer makes stuff like this trivial--then we should.

There are other good arguments. With a single fd, there is a single item to block on, which is mapped to a single queue of events. The single fd returns all watch events and also any potential out-of-band data. If every fd was a separate watch,

- There would be no way to get event ordering. Events on file foo and file bar would pop poll() on both fd's, but there would be no way to tell which happened first. A single queue trivially gives you ordering. Such ordering is crucial to existing applications such as Beagle. Imagine "mv a b ; mv b a" events without ordering.
- We'd have to maintain n fd's and n internal queues with state, versus just one. It is a lot messier in the kernel. A single, linear queue is the data structure that makes sense.
- User-space developers prefer the current API. The Beagle guys, for example, love it. Trust me, I asked. It is not a surprise: Who'd want to manage and block on 1000 fd's via select?
- No way to get out of band data.
- 1024 is still too low. ;-)

When you talk about designing a file change notification system that scales to 1000s of directories, juggling 1000s of fd's just does not seem the right interface. It is too heavy.

Additionally, it _is_ possible to more than one instance and juggle more than one queue and thus more than one associated fd. There need not be a one-fd-per-process mapping; it is one-fd-per-queue and a process can easily want more than one queue.

**Q:**

Why the system call approach?

**A:**

The poor user-space interface is the second biggest problem with dnotify. Signals are a terrible, terrible interface for file notification. Or for anything, for that matter. The ideal solution, from all perspectives, is a file descriptor-based one that allows basic file I/O and poll/select. Obtaining the fd and managing the watches could have been done either via a device file or a family of new system calls. We decided to implement a family of system calls because that is the preferred approach for new kernel interfaces. The only real difference was whether we wanted to use open(2) and ioctl(2) or a couple of new system calls. System calls beat ioctls.

## 3.33 ISO9660 Filesystem

Mount options that are the same as for msdos and vfat partitions.

| | |
|---|---|
| gid=nnn | All files in the partition will be in group nnn. |
| uid=nnn | All files in the partition will be owned by user id nnn. |
| umask=nnn | The permission mask (see umask(1)) for the partition. |

Mount options that are the same as vfat partitions. These are only useful when using discs encoded using Microsoft's Joliet extensions.

| | |
|---|---|
| iocharset=name | Character set to use for converting from Unicode to ASCII. Joliet filenames are stored in Unicode format, but Unix for the most part doesn't know how to deal with Unicode. There is also an option of doing UTF-8 translations with the utf8 option. |
| utf8 | Encode Unicode names in UTF-8 format. Default is no. |

Mount options unique to the isofs filesystem.

| | |
|---|---|
| block=512 | Set the block size for the disk to 512 bytes |
| block=1024 | Set the block size for the disk to 1024 bytes |
| block=2048 | Set the block size for the disk to 2048 bytes |
| check=relaxed | Matches filenames with different cases |
| check=strict | Matches only filenames with the exact same case |
| cruft | Try to handle badly formatted CDs. |
| map=off | Do not map non-Rock Ridge filenames to lower case |
| map=normal | Map non-Rock Ridge filenames to lower case |
| map=acorn | As map=normal but also apply Acorn extensions if present |
| mode=xxx | Sets the permissions on files to xxx unless Rock Ridge extensions set the permissions otherwise |
| dmode=xxx | Sets the permissions on directories to xxx unless Rock Ridge extensions set the permissions otherwise |
| overriderockperm | Set permissions on files and directories according to 'mode' and 'dmode' even though Rock Ridge extensions are present. |
| nojoliet | Ignore Joliet extensions if they are present. |
| norock | Ignore Rock Ridge extensions if they are present. |
| hide | Completely strip hidden files from the file system. |
| showassoc | Show files marked with the 'associated' bit |
| unhide | Deprecated; showing hidden files is now default; If given, it is a synonym for 'showassoc' which will recreate previous unhide behavior |
| session=x | Select number of session on multisession CD |
| sbsector=xxx | Session begins from sector xxx |

Recommended documents about ISO 9660 standard are located at:

- http://www.y-adagio.com/
- ftp://ftp.ecma.ch/ecma-st/Ecma-119.pdf

Quoting from the PDF ”This 2nd Edition of Standard ECMA-119 is technically identical with ISO 9660.”, so it is a valid and gratis substitute of the official ISO specification.

## 3.34 NILFS2

NILFS2 is a log-structured file system (LFS) supporting continuous snapshotting. In addition to versioning capability of the entire file system, users can even restore files mistakenly overwritten or destroyed just a few seconds ago. Since NILFS2 can keep consistency like conventional LFS, it achieves quick recovery after system crashes.

NILFS2 creates a number of checkpoints every few seconds or per synchronous write basis (unless there is no change). Users can select significant versions among continuously created checkpoints, and can change them into snapshots which will be preserved until they are changed back to checkpoints.

There is no limit on the number of snapshots until the volume gets full. Each snapshot is mountable as a read-only file system concurrently with its writable mount, and this feature is convenient for online backup.

The userland tools are included in nilfs-utils package, which is available from the following download page. At least ”mkfs.nilfs2”, ”mount.nilfs2”, ”umount.nilfs2”, and ”nilfs_cleanerd” (so called cleaner or garbage collector) are required. Details on the tools are described in the man pages included in the package.

**Project web page**
https://nilfs.sourceforge.io/

**Download page**
https://nilfs.sourceforge.io/en/download.html

**List info**
http://vger.kernel.org/vger-lists.html#linux-nilfs

### 3.34.1 Caveats

Features which NILFS2 does not support yet:

- atime
- extended attributes
- POSIX ACLs
- quotas
- fsck
- defragmentation

### 3.34.2 Mount options

NILFS2 supports the following mount options: (*) == default

| | |
|---|---|
| barrier(*) | This enables/disables the use of write barriers. This |
| nobarrier | requires an IO stack which can support barriers, and if nilfs gets an error on a barrier write, it will disable again with a warning. |
| errors=continue | Keep going on a filesystem error. |
| errors=remount-ro(*) | Remount the filesystem read-only on an error. |
| errors=panic | Panic and halt the machine if an error occurs. |
| cp=n | Specify the checkpoint-number of the snapshot to be mounted. Checkpoints and snapshots are listed by lscp user command. Only the checkpoints marked as snapshot are mountable with this option. Snapshot is read-only, so a read-only mount option must be specified together. |
| order=relaxed(*) | Apply relaxed order semantics that allows modified data blocks to be written to disk without making a checkpoint if no metadata update is going. This mode is equivalent to the ordered data mode of the ext3 filesystem except for the updates on data blocks still conserve atomicity. This will improve synchronous write performance for overwriting. |
| order=strict | Apply strict in-order semantics that preserves sequence of all file operations including overwriting of data blocks. That means, it is guaranteed that no overtaking of events occurs in the recovered file system after a crash. |
| norecovery | Disable recovery of the filesystem on mount. This disables every write access on the device for read-only mounts or snapshots. This option will fail for r/w mounts on an unclean volume. |
| discard | This enables/disables the use of discard/TRIM commands. |
| nodiscard(*) | The discard/TRIM commands are sent to the underlying block device when blocks are freed. This is useful for SSD devices and sparse/thinly-provisioned LUNs. |

### 3.34.3 Ioctls

There is some NILFS2 specific functionality which can be accessed by applications through the system call interfaces. The list of all NILFS2 specific ioctls are shown in the table below.

Table of NILFS2 specific ioctls:

| Ioctl | Description |
|---|---|
| NILFS_IOCTL_CHANGE_CPMODE | Change mode of given checkpoint between checkpoint and snapshot state. This ioctl is used in chcp and mkcp utilities. |
| NILFS_IOCTL_DELETE_CHECKPOINT | Remove checkpoint from NILFS2 file system. This ioctl is used in rmcp utility. |
| NILFS_IOCTL_GET_CPINFO | Return info about requested checkpoints. This ioctl is used in lscp utility and by nilfs_cleanerd daemon. |
| NILFS_IOCTL_GET_CPSTAT | Return checkpoints statistics. This ioctl is used by lscp, rmcp utilities and by nilfs_cleanerd daemon. |
| NILFS_IOCTL_GET_SUINFO | Return segment usage info about requested segments. This ioctl is used in lssu, nilfs_resize utilities and by nilfs_cleanerd daemon. |
| NILFS_IOCTL_SET_SUINFO | Modify segment usage info of requested segments. This ioctl is used by nilfs_cleanerd daemon to skip unnecessary cleaning operation of segments and reduce performance penalty or wear of flash device due to redundant move of in-use blocks. |
| NILFS_IOCTL_GET_SUSTAT | Return segment usage statistics. This ioctl is used in lssu, nilfs_resize utilities and by nilfs_cleanerd daemon. |
| NILFS_IOCTL_GET_VINFO | Return information on virtual block addresses. This ioctl is used by nilfs_cleanerd daemon. |
| NILFS_IOCTL_GET_BDESCS | Return information about descriptors of disk block numbers. This ioctl is used by nilfs_cleanerd daemon. |
| NILFS_IOCTL_CLEAN_SEGMENTS | Do garbage collection operation in the environment of requested parameters from userspace. This ioctl is used by nilfs_cleanerd daemon. |
| NILFS_IOCTL_SYNC | Make a checkpoint. This ioctl is used in mkcp utility. |
| NILFS_IOCTL_RESIZE | Resize NILFS2 volume. This ioctl is used by nilfs_resize utility. |
| NILFS_IOCTL_SET_ALLOC_RANGE | Define lower limit of segments in bytes and upper limit of segments in bytes. This ioctl is used by nilfs_resize utility. |

### 3.34.4 NILFS2 usage

To use nilfs2 as a local file system, simply:

```
# mkfs -t nilfs2 /dev/block_device
# mount -t nilfs2 /dev/block_device /dir
```

This will also invoke the cleaner through the mount helper program (mount.nilfs2).

Checkpoints and snapshots are managed by the following commands. Their manpages are included in the nilfs-utils package above.

| | |
|---|---|
| lscp | list checkpoints or snapshots. |
| mkcp | make a checkpoint or a snapshot. |
| chcp | change an existing checkpoint to a snapshot or vice versa. |
| rmcp | invalidate specified checkpoint(s). |

To mount a snapshot:

```
# mount -t nilfs2 -r -o cp=<cno> /dev/block_device /snap_dir
```

where <cno> is the checkpoint number of the snapshot.

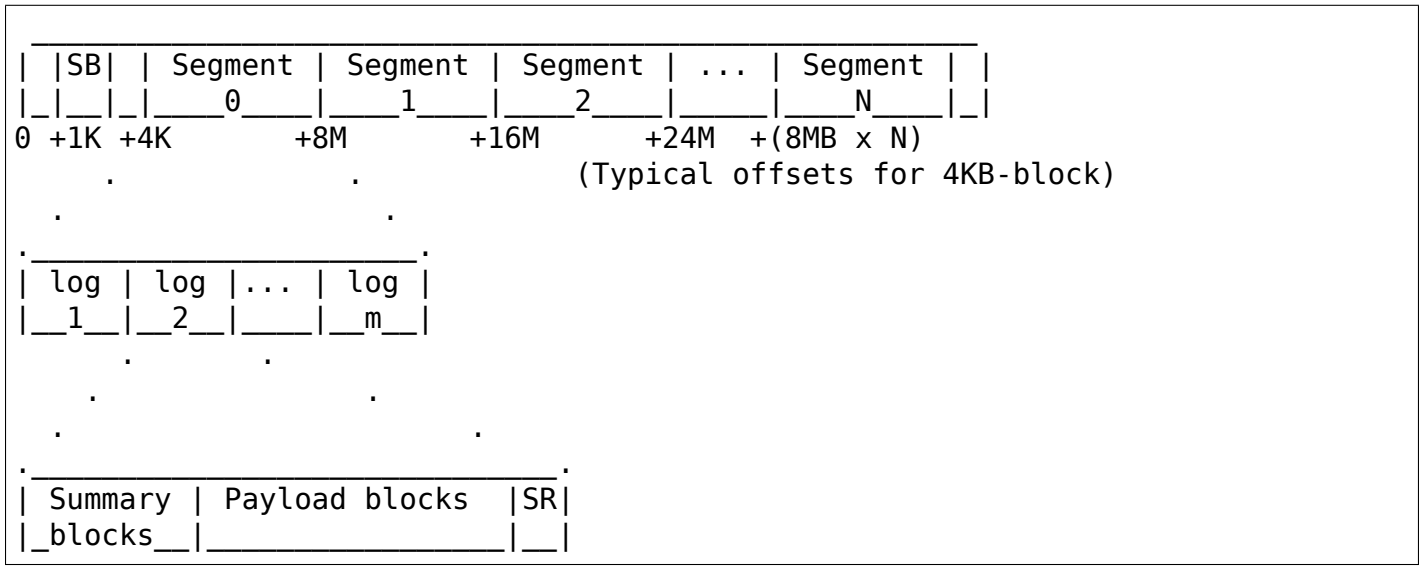
To unmount the NILFS2 mount point or snapshot, simply:

```
# umount /dir
```

Then, the cleaner daemon is automatically shut down by the umount helper program (umount.nilfs2).

### 3.34.5 Disk format

A nilfs2 volume is equally divided into a number of segments except for the super block (SB) and segment #0. A segment is the container of logs. Each log is composed of summary information blocks, payload blocks, and an optional super root block (SR):



The payload blocks are organized per file, and each file consists of data blocks and B-tree node blocks:

```
 |<---       File-A        --->|<---       File-B        --->|
_____________________________________________________________
 | Data blocks | B-tree blocks | Data blocks | B-tree blocks | ...
_|_____________|_______________|_____________|_______________|_
```

Since only the modified blocks are written in the log, it may have files without data blocks or B-tree node blocks.

The organization of the blocks is recorded in the summary information blocks, which contains a header structure (nilfs_segment_summary), per file structures (nilfs_finfo), and per block structures (nilfs_binfo):

```
 _____________________________________________________________________
| Summary | finfo | binfo | ... | binfo | finfo | binfo | ... | binfo |...
|_blocks__|___A___|_(A,1)_|_____|(A,Na)_|___B___|_(B,1)_|_____|(B,Nb)_|___
```

The logs include regular files, directory files, symbolic link files and several meta data files. The meta data files are the files used to maintain file system meta data. The current version of NILFS2 uses the following meta data files:

```
1) Inode file (ifile)             -- Stores on-disk inodes
2) Checkpoint file (cpfile)       -- Stores checkpoints
3) Segment usage file (sufile)    -- Stores allocation state of segments
4) Data address translation file  -- Maps virtual block numbers to usual
   (DAT)                             block numbers.  This file serves to
                                     make on-disk blocks relocatable.
```

The following figure shows a typical organization of the logs:

```
 ____________________________________________________________________
| Summary | regular file | file  | ... | ifile | cpfile | sufile | DAT |SR|
|_blocks__|_or_directory_|_______|_____|_______|________|________|_____|__|
```

To stride over segment boundaries, this sequence of files may be split into multiple logs. The sequence of logs that should be treated as logically one log, is delimited with flags marked in the segment summary. The recovery code of nilfs2 looks this boundary information to ensure atomicity of updates.

The super root block is inserted for every checkpoints. It includes three special inodes, inodes for the DAT, cpfile, and sufile. Inodes of regular files, directories, symlinks and other special files, are included in the ifile. The inode of ifile itself is included in the corresponding checkpoint entry in the cpfile. Thus, the hierarchy among NILFS2 files can be depicted as follows:

```
Super block (SB)
     |
     v
Super root block (the latest cno=xx)
     |-- DAT
     |-- sufile
     `-- cpfile
```

```
            |-- ifile (cno=c1)
            |-- ifile (cno=c2) ---- file (ino=i1)
            :        :            |-- file (ino=i2)
            `-- ifile (cno=xx)    |-- file (ino=i3)
                                  :        :
                                  `-- file (ino=yy)
                                    ( regular file, directory, or symlink )
```

For detail on the format of each file, please see nilfs2_ondisk.h located at include/uapi/linux directory.

There are no patents or other intellectual property that we protect with regard to the design of NILFS2. It is allowed to replicate the design in hopes that other operating systems could share (mount, read, write, etc.) data stored in this format.

## 3.35 NFS

### 3.35.1 NFSv4 client identifier

This document explains how the NFSv4 protocol identifies client instances in order to maintain file open and lock state during system restarts. A special identifier and principal are maintained on each client. These can be set by administrators, scripts provided by site administrators, or tools provided by Linux distributors.

There are risks if a client's NFSv4 identifier and its principal are not chosen carefully.

#### Introduction

The NFSv4 protocol uses ”lease-based file locking”. Leases help NFSv4 servers provide file lock guarantees and manage their resources.

Simply put, an NFSv4 server creates a lease for each NFSv4 client. The server collects each client's file open and lock state under the lease for that client.

The client is responsible for periodically renewing its leases. While a lease remains valid, the server holding that lease guarantees the file locks the client has created remain in place.

If a client stops renewing its lease (for example, if it crashes), the NFSv4 protocol allows the server to remove the client's open and lock state after a certain period of time. When a client restarts, it indicates to servers that open and lock state associated with its previous leases is no longer valid and can be destroyed immediately.

In addition, each NFSv4 server manages a persistent list of client leases. When the server restarts and clients attempt to recover their state, the server uses this list to distinguish amongst clients that held state before the server restarted and clients sending fresh OPEN and LOCK requests. This enables file locks to persist safely across server restarts.

### NFSv4 client identifiers

Each NFSv4 client presents an identifier to NFSv4 servers so that they can associate the client with its lease. Each client's identifier consists of two elements:

- co_ownerid: An arbitrary but fixed string.
- boot verifier: A 64-bit incarnation verifier that enables a server to distinguish successive boot epochs of the same client.

The NFSv4.0 specification refers to these two items as an ”nfs_client_id4”. The NFSv4.1 specification refers to these two items as a ”client_owner4”.

NFSv4 servers tie this identifier to the principal and security flavor that the client used when presenting it. Servers use this principal to authorize subsequent lease modification operations sent by the client. Effectively this principal is a third element of the identifier.

As part of the identity presented to servers, a good ”co_ownerid” string has several important properties:

- The ”co_ownerid” string identifies the client during reboot recovery, therefore the string is persistent across client reboots.
- The ”co_ownerid” string helps servers distinguish the client from others, therefore the string is globally unique. Note that there is no central authority that assigns ”co_ownerid” strings.
- Because it often appears on the network in the clear, the ”co_ownerid” string does not reveal private information about the client itself.
- The content of the ”co_ownerid” string is set and unchanging before the client attempts NFSv4 mounts after a restart.
- The NFSv4 protocol places a 1024-byte limit on the size of the ”co_ownerid” string.

### Protecting NFSv4 lease state

NFSv4 servers utilize the ”client_owner4” as described above to assign a unique lease to each client. Under this scheme, there are circumstances where clients can interfere with each other. This is referred to as ”lease stealing”.

If distinct clients present the same ”co_ownerid” string and use the same principal (for example, AUTH_SYS and UID 0), a server is unable to tell that the clients are not the same. Each distinct client presents a different boot verifier, so it appears to the server as if there is one client that is rebooting frequently. Neither client can maintain open or lock state in this scenario.

If distinct clients present the same ”co_ownerid” string and use distinct principals, the server is likely to allow the first client to operate normally but reject subsequent clients with the same ”co_ownerid” string.

If a client's ”co_ownerid” string or principal are not stable, state recovery after a server or client reboot is not guaranteed. If a client unexpectedly restarts but presents a different ”co_ownerid” string or principal to the server, the server orphans the client's previous open and lock state. This blocks access to locked files until the server removes the orphaned state.

If the server restarts and a client presents a changed ”co_ownerid” string or principal to the server, the server will not allow the client to reclaim its open and lock state, and may give those locks to other clients in the meantime. This is referred to as ”lock stealing”.

Lease stealing and lock stealing increase the potential for denial of service and in rare cases even data corruption.

### Selecting an appropriate client identifier

By default, the Linux NFSv4 client implementation constructs its ”co_ownerid” string starting with the words ”Linux NFS” followed by the client's UTS node name (the same node name, incidentally, that is used as the ”machine name” in an AUTH_SYS credential). In small deployments, this construction is usually adequate. Often, however, the node name by itself is not adequately unique, and can change unexpectedly. Problematic situations include:

- NFS-root (diskless) clients, where the local DHCP server (or equivalent) does not provide a unique host name.
- ”Containers” within a single Linux host. If each container has a separate network namespace, but does not use the UTS namespace to provide a unique host name, then there can be multiple NFS client instances with the same host name.
- Clients across multiple administrative domains that access a common NFS server. If hostnames are not assigned centrally then uniqueness cannot be guaranteed unless a domain name is included in the hostname.

Linux provides two mechanisms to add uniqueness to its ”co_ownerid” string:

**nfs.nfs4_unique_id**
: This module parameter can set an arbitrary uniquifier string via the kernel command line, or when the ”nfs” module is loaded.

**/sys/fs/nfs/net/nfs_client/identifier**
: This virtual file, available since Linux 5.3, is local to the network namespace in which it is accessed and so can provide distinction between network namespaces (containers) when the hostname remains uniform.

Note that this file is empty on name-space creation. If the container system has access to some sort of per-container identity then that uniquifier can be used. For example, a uniquifier might be formed at boot using the container's internal identifier:

**sha256sum /etc/machine-id | awk '{print $1}' **
: > /sys/fs/nfs/net/nfs_client/identifier

### Security considerations

The use of cryptographic security for lease management operations is strongly encouraged.

If NFS with Kerberos is not configured, a Linux NFSv4 client uses AUTH_SYS and UID 0 as the principal part of its client identity. This configuration is not only insecure, it increases the risk of lease and lock stealing. However, it might be the only choice for client configurations that have no local persistent storage. ”co_ownerid” string uniqueness and persistence is critical in this case.

When a Kerberos keytab is present on a Linux NFS client, the client attempts to use one of the principals in that keytab when identifying itself to servers. The ”sec=” mount option does not control this behavior. Alternately, a single-user client with a Kerberos principal can use that principal in place of the client's host principal.

Using Kerberos for this purpose enables the client and server to use the same lease for operations covered by all ”sec=” settings. Additionally, the Linux NFS client uses the RPCSEC_GSS security flavor with Kerberos and the integrity QOS to prevent in-transit modification of lease modification requests.

#### Additional notes

The Linux NFSv4 client establishes a single lease on each NFSv4 server it accesses. NFSv4 mounts from a Linux NFSv4 client of a particular server then share that lease.

Once a client establishes open and lock state, the NFSv4 protocol enables lease state to transition to other servers, following data that has been migrated. This hides data migration completely from running applications. The Linux NFSv4 client facilitates state migration by presenting the same ”client_owner4” to all servers it encounters.

### 3.35.2 See Also

- nfs(5)
- kerberos(7)
- RFC 7530 for the NFSv4.0 specification
- RFC 8881 for the NFSv4.1 specification.

### 3.35.3 Making Filesystems Exportable

#### Overview

All filesystem operations require a dentry (or two) as a starting point. Local applications have a reference-counted hold on suitable dentries via open file descriptors or cwd/root. However remote applications that access a filesystem via a remote filesystem protocol such as NFS may not be able to hold such a reference, and so need a different way to refer to a particular dentry. As the alternative form of reference needs to be stable across renames, truncates, and server-reboot (among other things, though these tend to be the most problematic), there is no simple answer like 'filename'.

The mechanism discussed here allows each filesystem implementation to specify how to generate an opaque (outside of the filesystem) byte string for any dentry, and how to find an appropriate dentry for any given opaque byte string. This byte string will be called a ”filehandle fragment” as it corresponds to part of an NFS filehandle.

A filesystem which supports the mapping between filehandle fragments and dentries will be termed ”exportable”.

### Dcache Issues

The dcache normally contains a proper prefix of any given filesystem tree. This means that if any filesystem object is in the dcache, then all of the ancestors of that filesystem object are also in the dcache. As normal access is by filename this prefix is created naturally and maintained easily (by each object maintaining a reference count on its parent).

However when objects are included into the dcache by interpreting a filehandle fragment, there is no automatic creation of a path prefix for the object. This leads to two related but distinct features of the dcache that are not needed for normal filesystem access.

1. The dcache must sometimes contain objects that are not part of the proper prefix. i.e that are not connected to the root.
2. The dcache must be prepared for a newly found (via ->lookup) directory to already have a (non-connected) dentry, and must be able to move that dentry into place (based on the parent and name in the ->lookup). This is particularly needed for directories as it is a dcache invariant that directories only have one dentry.

To implement these features, the dcache has:

a. A dentry flag DCACHE_DISCONNECTED which is set on any dentry that might not be part of the proper prefix. This is set when anonymous dentries are created, and cleared when a dentry is noticed to be a child of a dentry which is in the proper prefix. If the refcount on a dentry with this flag set becomes zero, the dentry is immediately discarded, rather than being kept in the dcache. If a dentry that is not already in the dcache is repeatedly accessed by filehandle (as NFSD might do), an new dentry will be a allocated for each access, and discarded at the end of the access.

   Note that such a dentry can acquire children, name, ancestors, etc. without losing DCACHE_DISCONNECTED - that flag is only cleared when subtree is successfully reconnected to root. Until then dentries in such subtree are retained only as long as there are references; refcount reaching zero means immediate eviction, same as for unhashed dentries. That guarantees that we won't need to hunt them down upon umount.

b. A primitive for creation of secondary roots - d_obtain_root(inode). Those do _not_ bear DCACHE_DISCONNECTED. They are placed on the per-superblock list (->s_roots), so they can be located at umount time for eviction purposes.

c. Helper routines to allocate anonymous dentries, and to help attach loose directory dentries at lookup time. They are:

   **d_obtain_alias(inode) will return a dentry for the given inode.**
   If the inode already has a dentry, one of those is returned.

   If it doesn't, a new anonymous (IS_ROOT and DCACHE_DISCONNECTED) dentry is allocated and attached.

   In the case of a directory, care is taken that only one dentry can ever be attached.

   **d_splice_alias(inode, dentry) will introduce a new dentry into the tree;**
   either the passed-in dentry or a preexisting alias for the given inode (such as an anonymous one created by d_obtain_alias), if appropriate. It returns NULL when the passed-in dentry is used, following the calling convention of ->lookup.

### Filesystem Issues

For a filesystem to be exportable it must:

1. provide the filehandle fragment routines described below.
2. make sure that d_splice_alias is used rather than d_add when ->lookup finds an inode for a given parent and name.

   If inode is NULL, d_splice_alias(inode, dentry) is equivalent to:

   ```
   d_add(dentry, inode), NULL
   ```

   Similarly, d_splice_alias(ERR_PTR(err), dentry) = ERR_PTR(err)

   Typically the ->lookup routine will simply end with a:

   ```
           return d_splice_alias(inode, dentry);
   }
   ```

A file system implementation declares that instances of the filesystem are exportable by setting the s_export_op field in the struct super_block. This field must point to a ”struct export_operations” struct which has the following members:

**encode_fh (mandatory)**
: Takes a dentry and creates a filehandle fragment which may later be used to find or create a dentry for the same object.

**fh_to_dentry (mandatory)**
: Given a filehandle fragment, this should find the implied object and create a dentry for it (possibly with d_obtain_alias).

**fh_to_parent (optional but strongly recommended)**
: Given a filehandle fragment, this should find the parent of the implied object and create a dentry for it (possibly with d_obtain_alias). May fail if the filehandle fragment is too small.

**get_parent (optional but strongly recommended)**
: When given a dentry for a directory, this should return a dentry for the parent. Quite possibly the parent dentry will have been allocated by d_alloc_anon. The default get_parent function just returns an error so any filehandle lookup that requires finding a parent will fail. ->lookup(”..”) is *not* used as a default as it can leave ”..” entries in the dcache which are too messy to work with.

**get_name (optional)**
: When given a parent dentry and a child dentry, this should find a name in the directory identified by the parent dentry, which leads to the object identified by the child dentry. If no get_name function is supplied, a default implementation is provided which uses vfs_readdir to find potential names, and matches inode numbers to find the correct match.

**flags**
: Some filesystems may need to be handled differently than others. The export_operations struct also includes a flags field that allows the filesystem to communicate such information to nfsd. See the Export Operations Flags section below for more explanation.

A filehandle fragment consists of an array of 1 or more 4byte words, together with a one byte "type". The decode_fh routine should not depend on the stated size that is passed to it. This size may be larger than the original filehandle generated by encode_fh, in which case it will have been padded with nuls. Rather, the encode_fh routine should choose a "type" which indicates the decode_fh how much of the filehandle is valid, and how it should be interpreted.

### Export Operations Flags

In addition to the operation vector pointers, struct export_operations also contains a "flags" field that allows the filesystem to communicate to nfsd that it may want to do things differently when dealing with it. The following flags are defined:

**EXPORT_OP_NOWCC - disable NFSv3 WCC attributes on this filesystem**
RFC 1813 recommends that servers always send weak cache consistency (WCC) data to the client after each operation. The server should atomically collect attributes about the inode, do an operation on it, and then collect the attributes afterward. This allows the client to skip issuing GETATTRs in some situations but means that the server is calling vfs_getattr for almost all RPCs. On some filesystems (particularly those that are clustered or networked) this is expensive and atomicity is difficult to guarantee. This flag indicates to nfsd that it should skip providing WCC attributes to the client in NFSv3 replies when doing operations on this filesystem. Consider enabling this on filesystems that have an expensive ->getattr inode operation, or when atomicity between pre and post operation attribute collection is impossible to guarantee.

**EXPORT_OP_NOSUBTREECHK - disallow subtree checking on this fs**
Many NFS operations deal with filehandles, which the server must then vet to ensure that they live inside of an exported tree. When the export consists of an entire filesystem, this is trivial. nfsd can just ensure that the filehandle live on the filesystem. When only part of a filesystem is exported however, then nfsd must walk the ancestors of the inode to ensure that it's within an exported subtree. This is an expensive operation and not all filesystems can support it properly. This flag exempts the filesystem from subtree checking and causes exportfs to get back an error if it tries to enable subtree checking on it.

**EXPORT_OP_CLOSE_BEFORE_UNLINK - always close cached files before unlinking**
On some exportable filesystems (such as NFS) unlinking a file that is still open can cause a fair bit of extra work. For instance, the NFS client will do a "sillyrename" to ensure that the file sticks around while it's still open. When reexporting, that open file is held by nfsd so we usually end up doing a sillyrename, and then immediately deleting the sillyrenamed file just afterward when the link count actually goes to zero. Sometimes this delete can race with other operations (for instance an rmdir of the parent directory). This flag causes nfsd to close any open files for this inode _before_ calling into the vfs to do an unlink or a rename that would replace an existing file.

**EXPORT_OP_REMOTE_FS - Backing storage for this filesystem is remote**
PF_LOCAL_THROTTLE exists for loopback NFSD, where a thread needs to write to one bdi (the final bdi) in order to free up writes queued to another bdi (the client bdi). Such threads get a private balance of dirty pages so that dirty pages for the client bdi do not imact the daemon writing to the final bdi. For filesystems

whose durable storage is not local (such as exported NFS filesystems), this constraint has negative consequences. EXPORT_OP_REMOTE_FS enables an export to disable writeback throttling.

**EXPORT_OP_NOATOMIC_ATTR - Filesystem does not update attributes atomically**
EXPORT_OP_NOATOMIC_ATTR indicates that the exported filesystem cannot provide the semantics required by the ”atomic” boolean in NFSv4's change_info4. This boolean indicates to a client whether the returned before and after change attributes were obtained atomically with the respect to the requested metadata operation (UNLINK, OPEN/CREATE, MKDIR, etc).

**EXPORT_OP_FLUSH_ON_CLOSE - Filesystem flushes file data on close(2)**
On most filesystems, inodes can remain under writeback after the file is closed. NFSD relies on client activity or local flusher threads to handle writeback. Certain filesystems, such as NFS, flush all of an inode's dirty data on last close. Exports that behave this way should set EXPORT_OP_FLUSH_ON_CLOSE so that NFSD knows to skip waiting for writeback when closing such files.

**EXPORT_OP_ASYNC_LOCK - Indicates a capable filesystem to do async lock**
requests from lockd. Only set EXPORT_OP_ASYNC_LOCK if the filesystem has it's own ->lock() functionality as core *posix_lock_file()* implementation has no async lock request handling yet. For more information about how to indicate an async lock request from a ->lock() file_operations struct, see fs/locks.c and comment for the function *vfs_lock_file()*.

### 3.35.4 Reference counting in pnfs

The are several inter-related caches. We have layouts which can reference multiple devices, each of which can reference multiple data servers. Each data server can be referenced by multiple devices. Each device can be referenced by multiple layouts. To keep all of this straight, we need to reference count.

#### struct pnfs_layout_hdr

The on-the-wire command LAYOUTGET corresponds to struct pnfs_layout_segment, usually referred to by the variable name lseg. Each nfs_inode may hold a pointer to a cache of these layout segments in nfsi->layout, of type struct pnfs_layout_hdr.

We reference the header for the inode pointing to it, across each outstanding RPC call that references it (LAYOUTGET, LAYOUTRETURN, LAYOUTCOMMIT), and for each lseg held within.

Each header is also (when non-empty) put on a list associated with struct nfs_client (cl_layouts). Being put on this list does not bump the reference count, as the layout is kept around by the lseg that keeps it in the list.

#### deviceid_cache

lsegs reference device ids, which are resolved per nfs_client and layout driver type. The device ids are held in a RCU cache (struct nfs4_deviceid_cache). The cache itself is referenced across each mount. The entries (struct nfs4_deviceid) themselves are held across the lifetime of each lseg referencing them.

RCU is used because the deviceid is basically a write once, read many data structure. The hlist size of 32 buckets needs better justification, but seems reasonable given that we can have multiple deviceid's per filesystem, and multiple filesystems per nfs_client.

The hash code is copied from the nfsd code base. A discussion of hashing and variations of this algorithm can be found here.

#### data server cache

file driver devices refer to data servers, which are kept in a module level cache. Its reference is held over the lifetime of the deviceid pointing to it.

#### lseg

lseg maintains an extra reference corresponding to the NFS_LSEG_VALID bit which holds it in the pnfs_layout_hdr's list. When the final lseg is removed from the pnfs_layout_hdr's list, the NFS_LAYOUT_DESTROYED bit is set, preventing any new lsegs from being added.

#### layout drivers

PNFS utilizes what is called layout drivers. The STD defines 4 basic layout types: ”files”, ”objects”, ”blocks”, and ”flexfiles”. For each of these types there is a layout-driver with a common function-vectors table which are called by the nfs-client pnfs-core to implement the different layout types.

Files-layout-driver code is in: fs/nfs/filelayout/.. directory Blocks-layout-driver code is in: fs/nfs/blocklayout/.. directory Flexfiles-layout-driver code is in: fs/nfs/flexfilelayout/.. directory

#### blocks-layout setup

TODO: Document the setup needs of the blocks layout driver

### 3.35.5 RPC Cache

This document gives a brief introduction to the caching mechanisms in the sunrpc layer that is used, in particular, for NFS authentication.

### Caches

The caching replaces the old exports table and allows for a wide variety of values to be caches.

There are a number of caches that are similar in structure though quite possibly very different in content and use. There is a corpus of common code for managing these caches.

Examples of caches that are likely to be needed are:

- mapping from IP address to client name
- mapping from client name and filesystem to export options
- mapping from UID to list of GIDs, to work around NFS's limitation of 16 gids.
- mappings between local UID/GID and remote UID/GID for sites that do not have uniform uid assignment
- mapping from network identify to public key for crypto authentication.

The common code handles such things as:

- general cache lookup with correct locking
- supporting 'NEGATIVE' as well as positive entries
- allowing an EXPIRED time on cache items, and removing items after they expire, and are no longer in-use.
- making requests to user-space to fill in cache entries
- allowing user-space to directly set entries in the cache
- delaying RPC requests that depend on as-yet incomplete cache entries, and replaying those requests when the cache entry is complete.
- clean out old entries as they expire.

### Creating a Cache

- A cache needs a datum to store. This is in the form of a structure definition that must contain a struct cache_head as an element, usually the first. It will also contain a key and some content. Each cache element is reference counted and contains expiry and update times for use in cache management.
- A cache needs a ”cache_detail” structure that describes the cache. This stores the hash table, some parameters for cache management, and some operations detailing how to work with particular cache items.

  The operations are:

  **struct cache_head *alloc(void)**
  : This simply allocates appropriate memory and returns a pointer to the cache_detail embedded within the structure

  **void cache_put(struct kref *)**
  : This is called when the last reference to an item is dropped. The pointer passed is to the 'ref' field in the cache_head. cache_put should release any references create by 'cache_init' and, if CACHE_VALID is set, any references

created by cache_update. It should then release the memory allocated by 'alloc'.

**int match(struct cache_head *orig, struct cache_head *new)**
test if the keys in the two structures match. Return 1 if they do, 0 if they don't.

**void init(struct cache_head *orig, struct cache_head *new)**
Set the 'key' fields in 'new' from 'orig'. This may include taking references to shared objects.

**void update(struct cache_head *orig, struct cache_head *new)**
Set the 'content' fields in 'new' from 'orig'.

**int cache_show(struct seq_file *m, struct cache_detail *cd, struct cache_head *h)**
Optional. Used to provide a /proc file that lists the contents of a cache. This should show one item, usually on just one line.

**int cache_request(struct cache_detail *cd, struct cache_head *h, char **bpp, int *blen)**
Format a request to be send to user-space for an item to be instantiated. *bpp is a buffer of size *blen. bpp should be moved forward over the encoded message, and *blen should be reduced to show how much free space remains. Return 0 on success or <0 if not enough room or other problem.

**int cache_parse(struct cache_detail *cd, char *buf, int len)**
A message from user space has arrived to fill out a cache entry. It is in 'buf' of length 'len'. cache_parse should parse this, find the item in the cache with sunrpc_cache_lookup_rcu, and update the item with sunrpc_cache_update.

- A cache needs to be registered using cache_register(). This includes it on a list of caches that will be regularly cleaned to discard old data.

### Using a cache

To find a value in a cache, call sunrpc_cache_lookup_rcu passing a pointer to the cache_head in a sample item with the 'key' fields filled in. This will be passed to ->match to identify the target entry. If no entry is found, a new entry will be create, added to the cache, and marked as not containing valid data.

The item returned is typically passed to cache_check which will check if the data is valid, and may initiate an up-call to get fresh data. cache_check will return -ENOENT in the entry is negative or if an up call is needed but not possible, -EAGAIN if an upcall is pending, or 0 if the data is valid;

cache_check can be passed a ”struct cache_req*”. This structure is typically embedded in the actual request and can be used to create a deferred copy of the request (struct cache_deferred_req). This is done when the found cache item is not uptodate, but the is reason to believe that userspace might provide information soon. When the cache item does become valid, the deferred copy of the request will be revisited (->revisit). It is expected that this method will reschedule the request for processing.

The value returned by sunrpc_cache_lookup_rcu can also be passed to sunrpc_cache_update to set the content for the item. A second item is passed which should hold the content. If the item found by _lookup has valid data, then it is discarded and a new item is created. This saves any user of an item from worrying about content changing while it is being inspected. If

the item found by _lookup does not contain valid data, then the content is copied across and CACHE_VALID is set.

### Populating a cache

Each cache has a name, and when the cache is registered, a directory with that name is created in /proc/net/rpc

This directory contains a file called 'channel' which is a channel for communicating between kernel and user for populating the cache. This directory may later contain other files of interacting with the cache.

The 'channel' works a bit like a datagram socket. Each 'write' is passed as a whole to the cache for parsing and interpretation. Each cache can treat the write requests differently, but it is expected that a message written will contain:

- a key
- an expiry time
- a content.

with the intention that an item in the cache with the give key should be create or updated to have the given content, and the expiry time should be set on that item.

Reading from a channel is a bit more interesting. When a cache lookup fails, or when it succeeds but finds an entry that may soon expire, a request is lodged for that cache item to be updated by user-space. These requests appear in the channel file.

Successive reads will return successive requests. If there are no more requests to return, read will return EOF, but a select or poll for read will block waiting for another request to be added.

Thus a user-space helper is likely to:

```
open the channel.
  select for readable
  read a request
  write a response
loop.
```

If it dies and needs to be restarted, any requests that have not been answered will still appear in the file and will be read by the new instance of the helper.

Each cache should define a ”cache_parse” method which takes a message written from user-space and processes it. It should return an error (which propagates back to the write syscall) or 0.

Each cache should also define a ”cache_request” method which takes a cache item and encodes a request into the buffer provided.

---

**Note:** If a cache has no active readers on the channel, and has had not active readers for more than 60 seconds, further requests will not be added to the channel but instead all lookups that do not find a valid entry will fail. This is partly for backward compatibility: The previous nfs exports table was deemed to be authoritative and a failed lookup meant a definite 'no'.

---

#### request/response format

While each cache is free to use its own format for requests and responses over channel, the following is recommended as appropriate and support routines are available to help: Each request or response record should be printable ASCII with precisely one newline character which should be at the end. Fields within the record should be separated by spaces, normally one. If spaces, newlines, or nul characters are needed in a field they much be quoted. two mechanisms are available:

- If a field begins 'x' then it must contain an even number of hex digits, and pairs of these digits provide the bytes in the field.
- otherwise a in the field must be followed by 3 octal digits which give the code for a byte. Other characters are treated as them selves. At the very least, space, newline, nul, and '' must be quoted in this way.

### 3.35.6 rpcsec_gss support for kernel RPC servers

This document gives references to the standards and protocols used to implement RPCGSS authentication in kernel RPC servers such as the NFS server and the NFS client's NFSv4.0 callback server. (But note that NFSv4.1 and higher don't require the client to act as a server for the purposes of authentication.)

RPCGSS is specified in a few IETF documents:

- RFC2203 v1: https://tools.ietf.org/rfc/rfc2203.txt
- RFC5403 v2: https://tools.ietf.org/rfc/rfc5403.txt

There is a third version that we don't currently implement:

- RFC7861 v3: https://tools.ietf.org/rfc/rfc7861.txt

#### Background

The RPCGSS Authentication method describes a way to perform GSSAPI Authentication for NFS. Although GSSAPI is itself completely mechanism agnostic, in many cases only the KRB5 mechanism is supported by NFS implementations.

The Linux kernel, at the moment, supports only the KRB5 mechanism, and depends on GSSAPI extensions that are KRB5 specific.

GSSAPI is a complex library, and implementing it completely in kernel is unwarranted. However GSSAPI operations are fundamentally separable in 2 parts:

- initial context establishment
- integrity/privacy protection (signing and encrypting of individual packets)

The former is more complex and policy-independent, but less performance-sensitive. The latter is simpler and needs to be very fast.

Therefore, we perform per-packet integrity and privacy protection in the kernel, but leave the initial context establishment to userspace. We need upcalls to request userspace to perform context establishment.

#### NFS Server Legacy Upcall Mechanism

The classic upcall mechanism uses a custom text based upcall mechanism to talk to a custom daemon called rpc.svcgssd that is provide by the nfs-utils package.

This upcall mechanism has 2 limitations:

A) It can handle tokens that are no bigger than 2KiB

In some Kerberos deployment GSSAPI tokens can be quite big, up and beyond 64KiB in size due to various authorization extensions attacked to the Kerberos tickets, that needs to be sent through the GSS layer in order to perform context establishment.

B) It does not properly handle creds where the user is member of more than a few thousand groups (the current hard limit in the kernel is 65K groups) due to limitation on the size of the buffer that can be send back to the kernel (4KiB).

#### NFS Server New RPC Upcall Mechanism

The newer upcall mechanism uses RPC over a unix socket to a daemon called gss-proxy, implemented by a userspace program called Gssproxy.

The gss_proxy RPC protocol is currently documented here.

This upcall mechanism uses the kernel rpc client and connects to the gssproxy userspace program over a regular unix socket. The gssproxy protocol does not suffer from the size limitations of the legacy protocol.

#### Negotiating Upcall Mechanisms

To provide backward compatibility, the kernel defaults to using the legacy mechanism. To switch to the new mechanism, gss-proxy must bind to /var/run/gssproxy.sock and then write ”1” to /proc/net/rpc/use-gss-proxy. If gss-proxy dies, it must repeat both steps.

Once the upcall mechanism is chosen, it cannot be changed. To prevent locking into the legacy mechanisms, the above steps must be performed before starting nfsd. Whoever starts nfsd can guarantee this by reading from /proc/net/rpc/use-gss-proxy and checking that it contains a ”1”--the read will block until gss-proxy has done its write to the file.

### 3.35.7 NFSv4.1 Server Implementation

Server support for minorversion 1 can be controlled using the /proc/fs/nfsd/versions control file. The string output returned by reading this file will contain either ”+4.1” or ”-4.1” correspondingly.

Currently, server support for minorversion 1 is enabled by default. It can be disabled at run time by writing the string ”-4.1” to the /proc/fs/nfsd/versions control file. Note that to write this control file, the nfsd service must be taken down. You can use rpc.nfsd for this; see rpc.nfsd(8).

(Warning: older servers will interpret ”+4.1” and ”-4.1” as ”+4” and ”-4”, respectively. Therefore, code meant to work on both new and old kernels must turn 4.1 on or off *before* turning support for version 4 on or off; rpc.nfsd does this correctly.)

The NFSv4 minorversion 1 (NFSv4.1) implementation in nfsd is based on RFC 5661.

From the many new features in NFSv4.1 the current implementation focuses on the mandatory-to-implement NFSv4.1 Sessions, providing ”exactly once” semantics and better control and throttling of the resources allocated for each client.

The table below, taken from the NFSv4.1 document, lists the operations that are mandatory to implement (REQ), optional (OPT), and NFSv4.0 operations that are required not to implement (MNI) in minor version 1. The first column indicates the operations that are not supported yet by the linux server implementation.

The OPTIONAL features identified and their abbreviations are as follows:

- **pNFS** Parallel NFS
- **FDELG** File Delegations
- **DDELG** Directory Delegations

The following abbreviations indicate the linux server implementation status.

- **I** Implemented NFSv4.1 operations.
- **NS** Not Supported.
- **NS*** Unimplemented optional feature.

### Operations

| Implementation status | Operation | REQ,REC, OPT or NMI | Feature (REQ, REC or C |
|---|---|---|---|
| | ACCESS | REQ | |
| I | BACKCHANNEL_CTL | REQ | |
| I | BIND_CONN_TO_SESSION | REQ | |
| | CLOSE | REQ | |
| | COMMIT | REQ | |
| | CREATE | REQ | |
| I | CREATE_SESSION | REQ | |
| NS* | DELEGPURGE | OPT | FDELG (REQ) |
| | DELEGRETURN | OPT | FDELG, |
| | | | DDELG, pNFS |
| | | | (REQ) |
| I | DESTROY_CLIENTID | REQ | |
| I | DESTROY_SESSION | REQ | |
| I | EXCHANGE_ID | REQ | |
| I | FREE_STATEID | REQ | |
| | GETATTR | REQ | |
| I | GETDEVICEINFO | OPT | pNFS (REQ) |
| NS* | GETDEVICELIST | OPT | pNFS (OPT) |
| | GETFH | REQ | |
| NS* | GET_DIR_DELEGATION | OPT | DDELG (REQ) |
| I | LAYOUTCOMMIT | OPT | pNFS (REQ) |
| I | LAYOUTGET | OPT | pNFS (REQ) |
| I | LAYOUTRETURN | OPT | pNFS (REQ) |
| | LINK | OPT | |
| | LOCK | REQ | |

CO

Table 5 – continued from previous page

| Implementation status | Operation | REQ,REC, OPT or NMI | Feature (REQ, REC or C |
|---|---|---|---|
| | LOCKT | REQ | |
| | LOCKU | REQ | |
| | LOOKUP | REQ | |
| | LOOKUPP | REQ | |
| | NVERIFY | REQ | |
| | OPEN | REQ | |
| NS* | OPENATTR | OPT | |
| | OPEN_CONFIRM | MNI | |
| | OPEN_DOWNGRADE | REQ | |
| | PUTFH | REQ | |
| | PUTPUBFH | REQ | |
| | PUTROOTFH | REQ | |
| | READ | REQ | |
| | READDIR | REQ | |
| | READLINK | OPT | |
| | RECLAIM_COMPLETE | REQ | |
| | RELEASE_LOCKOWNER | MNI | |
| | REMOVE | REQ | |
| | RENAME | REQ | |
| | RENEW | MNI | |
| | RESTOREFH | REQ | |
| | SAVEFH | REQ | |
| | SECINFO | REQ | |
| I | SECINFO_NO_NAME | REC | pNFS files |
| | | | layout (REQ) |
| I | SEQUENCE | REQ | |
| | SETATTR | REQ | |
| | SETCLIENTID | MNI | |
| | SETCLIENTID_CONFIRM | MNI | |
| NS | SET_SSV | REQ | |
| I | TEST_STATEID | REQ | |
| | VERIFY | REQ | |
| NS* | WANT_DELEGATION | OPT | FDELG (OPT) |
| | WRITE | REQ | |

### Callback Operations

| Implementation status | Operation | REQ,REC, OPT or NMI | Feature (REQ, REC or OPT) | Definition |
|---|---|---|---|---|
| | CB_GETATTR | OPT | FDELG (REQ) | Section 20.1 |
| I | CB_LAYOUTRECALL | OPT | pNFS (REQ) | Section 20.3 |
| NS* | CB_NOTIFY | OPT | DDELG (REQ) | Section 20.4 |
| NS* | CB_NOTIFY_DEVICEID | OPT | pNFS (OPT) | Section 20.12 |
| NS* | CB_NOTIFY_LOCK | OPT | | Section 20.11 |
| NS* | CB_PUSH_DELEG | OPT | FDELG (OPT) | Section 20.5 |
| | CB_RECALL | OPT | FDELG, | Section 20.2 |
| | | | DDELG, pNFS | |
| | | | (REQ) | |
| NS* | CB_RECALL_ANY | OPT | FDELG, | Section 20.6 |
| | | | DDELG, pNFS | |
| | | | (REQ) | |
| NS | CB_RECALL_SLOT | REQ | | Section 20.8 |
| NS* | CB_RECALLABLE_OBJ_AVAIL | OPT | DDELG, pNFS | Section 20.7 |
| | | | (REQ) | |
| I | CB_SEQUENCE | OPT | FDELG, | Section 20.9 |
| | | | DDELG, pNFS | |
| | | | (REQ) | |
| NS* | CB_WANTS_CANCELLED | OPT | FDELG, | Section 20.10 |
| | | | DDELG, pNFS | |
| | | | (REQ) | |

### Implementation notes:

**SSV:**
The spec claims this is mandatory, but we don't actually know of any implementations, so we're ignoring it for now. The server returns NFS4ERR_ENCR_ALG_UNSUPP on EXCHANGE_ID, which should be future-proof.

**GSS on the backchannel:**
Again, theoretically required but not widely implemented (in particular, the current Linux client doesn't request it). We return NFS4ERR_ENCR_ALG_UNSUPP on CREATE_SESSION.

**DELEGPURGE:**
mandatory only for servers that support CLAIM_DELEGATE_PREV and/or CLAIM_DELEG_PREV_FH (which allows clients to keep delegations that persist across client reboots). Thus we need not implement this for now.

**EXCHANGE_ID:**
implementation ids are ignored

**CREATE_SESSION:**
backchannel attributes are ignored

**SEQUENCE:**
no support for dynamic slot table renegotiation (optional)

**Nonstandard compound limitations:**
No support for a sessions fore channel RPC compound that requires both a ca_maxrequestsize request and a ca_maxresponsesize reply, so we may fail to live up to the promise we made in CREATE_SESSION fore channel negotiation.

See also http://wiki.linux-nfs.org/wiki/index.php/Server_4.0_and_4.1_issues.

### 3.35.8 Kernel NFS Server Statistics

**Authors**
Greg Banks <gnb@sgi.com> - 26 Mar 2009

This document describes the format and semantics of the statistics which the kernel NFS server makes available to userspace. These statistics are available in several text form pseudo files, each of which is described separately below.

In most cases you don't need to know these formats, as the nfsstat(8) program from the nfs-utils distribution provides a helpful command-line interface for extracting and printing them.

All the files described here are formatted as a sequence of text lines, separated by newline 'n' characters. Lines beginning with a hash '#' character are comments intended for humans and should be ignored by parsing routines. All other lines contain a sequence of fields separated by whitespace.

#### /proc/fs/nfsd/pool_stats

This file is available in kernels from 2.6.30 onwards, if the /proc/fs/nfsd filesystem is mounted (it almost always should be).

The first line is a comment which describes the fields present in all the other lines. The other lines present the following data as a sequence of unsigned decimal numeric fields. One line is shown for each NFS thread pool.

All counters are 64 bits wide and wrap naturally. There is no way to zero these counters, instead applications should do their own rate conversion.

**pool**
The id number of the NFS thread pool to which this line applies. This number does not change.

Thread pool ids are a contiguous set of small integers starting at zero. The maximum value depends on the thread pool mode, but currently cannot be larger than the number of CPUs in the system. Note that in the default case there will be a single thread pool which contains all the nfsd threads and all the CPUs in the system, and thus this file will have a single line with a pool id of ”0”.

**packets-arrived**
Counts how many NFS packets have arrived. More precisely, this is the number of times that the network stack has notified the sunrpc server layer that new data may be available on a transport (e.g. an NFS or UDP socket or an NFS/RDMA endpoint).

Depending on the NFS workload patterns and various network stack effects (such as Large Receive Offload) which can combine packets on the wire, this may be either more or less than the number of NFS calls received (which statistic is available elsewhere). However

this is a more accurate and less workload-dependent measure of how much CPU load is being placed on the sunrpc server layer due to NFS network traffic.

**sockets-enqueued**
Counts how many times an NFS transport is enqueued to wait for an nfsd thread to service it, i.e. no nfsd thread was considered available.

The circumstance this statistic tracks indicates that there was NFS network-facing work to be done but it couldn't be done immediately, thus introducing a small delay in servicing NFS calls. The ideal rate of change for this counter is zero; significantly non-zero values may indicate a performance limitation.

This can happen because there are too few nfsd threads in the thread pool for the NFS workload (the workload is thread-limited), in which case configuring more nfsd threads will probably improve the performance of the NFS workload.

**threads-woken**
Counts how many times an idle nfsd thread is woken to try to receive some data from an NFS transport.

This statistic tracks the circumstance where incoming network-facing NFS work is being handled quickly, which is a good thing. The ideal rate of change for this counter will be close to but less than the rate of change of the packets-arrived counter.

**threads-timedout**
Counts how many times an nfsd thread triggered an idle timeout, i.e. was not woken to handle any incoming network packets for some time.

This statistic counts a circumstance where there are more nfsd threads configured than can be used by the NFS workload. This is a clue that the number of nfsd threads can be reduced without affecting performance. Unfortunately, it's only a clue and not a strong indication, for a couple of reasons:

- Currently the rate at which the counter is incremented is quite slow; the idle timeout is 60 minutes. Unless the NFS workload remains constant for hours at a time, this counter is unlikely to be providing information that is still useful.
- It is usually a wise policy to provide some slack, i.e. configure a few more nfsds than are currently needed, to allow for future spikes in load.

Note that incoming packets on NFS transports will be dealt with in one of three ways. An nfsd thread can be woken (threads-woken counts this case), or the transport can be enqueued for later attention (sockets-enqueued counts this case), or the packet can be temporarily deferred because the transport is currently being used by an nfsd thread. This last case is not very interesting and is not explicitly counted, but can be inferred from the other counters thus:

```
packets-deferred = packets-arrived - ( sockets-enqueued + threads-woken )
```

#### More

Descriptions of the other statistics file should go here.

### 3.35.9 Reexporting NFS filesystems

#### Overview

It is possible to reexport an NFS filesystem over NFS. However, this feature comes with a number of limitations. Before trying it, we recommend some careful research to determine whether it will work for your purposes.

A discussion of current known limitations follows.

#### ”fsid=” required, crossmnt broken

We require the ”fsid=” export option on any reexport of an NFS filesystem. You can use ”uuidgen -r” to generate a unique argument.

The ”crossmnt” export does not propagate ”fsid=”, so it will not allow traversing into further nfs filesystems; if you wish to export nfs filesystems mounted under the exported filesystem, you'll need to export them explicitly, assigning each its own unique ”fsid= option.

#### Reboot recovery

The NFS protocol's normal reboot recovery mechanisms don't work for the case when the reexport server reboots. Clients will lose any locks they held before the reboot, and further IO will result in errors. Closing and reopening files should clear the errors.

#### Filehandle limits

If the original server uses an X byte filehandle for a given object, the reexport server's filehandle for the reexported object will be X+22 bytes, rounded up to the nearest multiple of four bytes.

The result must fit into the RFC-mandated filehandle size limits:

| | |
|---|---|
| NFSv2 | 32 bytes |
| NFSv3 | 64 bytes |
| NFSv4 | 128 bytes |

So, for example, you will only be able to reexport a filesystem over NFSv2 if the original server gives you filehandles that fit in 10 bytes--which is unlikely.

In general there's no way to know the maximum filehandle size given out by an NFS server without asking the server vendor.

But the following table gives a few examples. The first column is the typical length of the filehandle from a Linux server exporting the given filesystem, the second is the length after that nfs export is reexported by another Linux host:

| | filehandle length | after reexport |
|---|---|---|
| ext4: | 28 bytes | 52 bytes |
| xfs: | 32 bytes | 56 bytes |
| btrfs: | 40 bytes | 64 bytes |

All will therefore fit in an NFSv3 or NFSv4 filehandle after reexport, but none are reexportable over NFSv2.

Linux server filehandles are a bit more complicated than this, though; for example:

- The (non-default) ”subtreecheck” export option generally requires another 4 to 8 bytes in the filehandle.
- If you export a subdirectory of a filesystem (instead of exporting the filesystem root), that also usually adds 4 to 8 bytes.
- If you export over NFSv2, knfsd usually uses a shorter filesystem identifier that saves 8 bytes.
- The root directory of an export uses a filehandle that is shorter.

As you can see, the 128-byte NFSv4 filehandle is large enough that you're unlikely to have trouble using NFSv4 to reexport any filesystem exported from a Linux server. In general, if the original server is something that also supports NFSv3, you're *probably* OK. Re-exporting over NFSv3 may be dicier, and reexporting over NFSv2 will probably never work.

For more details of Linux filehandle structure, the best reference is the source code and comments; see in particular:

- include/linux/exportfs.h:enum fid_type
- include/uapi/linux/nfsd/nfsfh.h:struct nfs_fhbase_new
- fs/nfsd/nfsfh.c:set_version_and_fsid_type
- fs/nfs/export.c:nfs_encode_fh

### Open DENY bits ignored

NFS since NFSv4 supports ALLOW and DENY bits taken from Windows, which allow you, for example, to open a file in a mode which forbids other read opens or write opens. The Linux client doesn't use them, and the server's support has always been incomplete: they are enforced only against other NFS users, not against processes accessing the exported filesystem locally. A reexport server will also not pass them along to the original server, so they will not be enforced between clients of different reexport servers.

# 3.36 The Linux NTFS filesystem driver

## 3.36.1 Overview

Linux-NTFS comes with a number of user-space programs known as ntfsprogs. These include mkntfs, a full-featured ntfs filesystem format utility, ntfsundelete used for recovering files that were unintentionally deleted from an NTFS volume and ntfsresize which is used to resize an NTFS partition. See the web site for more information.

To mount an NTFS 1.2/3.x (Windows NT4/2000/XP/2003) volume, use the file system type 'ntfs'. The driver currently supports read-only mode (with no fault-tolerance, encryption or journalling) and very limited, but safe, write support.

For fault tolerance and raid support (i.e. volume and stripe sets), you can use the kernel's Software RAID / MD driver. See section "Using Software RAID with NTFS" for details.

## 3.36.2 Web site

There is plenty of additional information on the linux-ntfs web site at http://www.linux-ntfs.org/

The web site has a lot of additional information, such as a comprehensive FAQ, documentation on the NTFS on-disk format, information on the Linux-NTFS userspace utilities, etc.

## 3.36.3 Features

- This is a complete rewrite of the NTFS driver that used to be in the 2.4 and earlier kernels. This new driver implements NTFS read support and is functionally equivalent to the old ntfs driver and it also implements limited write support. The biggest limitation at present is that files/directories cannot be created or deleted. See below for the list of write features that are so far supported. Another limitation is that writing to compressed files is not implemented at all. Also, neither read nor write access to encrypted files is so far implemented.
- The new driver has full support for sparse files on NTFS 3.x volumes which the old driver isn't happy with.
- The new driver supports execution of binaries due to mmap() now being supported.
- The new driver supports loopback mounting of files on NTFS which is used by some Linux distributions to enable the user to run Linux from an NTFS partition by creating a large file while in Windows and then loopback mounting the file while in Linux and creating a Linux filesystem on it that is used to install Linux on it.
- A comparison of the two drivers using:

```
time find . -type f -exec md5sum "{}" \;
```

  run three times in sequence with each driver (after a reboot) on a 1.4GiB NTFS partition, showed the new driver to be 20% faster in total time elapsed (from 9:43 minutes on average down to 7:53). The time spent in user space was unchanged but the time spent in the kernel was decreased by a factor of 2.5 (from 85 CPU seconds down to 33).

- The driver does not support short file names in general. For backwards compatibility, we implement access to files using their short file names if they exist. The driver will not create short file names however, and a rename will discard any existing short file name.
- The new driver supports exporting of mounted NTFS volumes via NFS.
- The new driver supports async io (aio).
- The new driver supports fsync(2), fdatasync(2), and msync(2).
- The new driver supports readv(2) and writev(2).
- The new driver supports access time updates (including mtime and ctime).
- The new driver supports truncate(2) and open(2) with O_TRUNC. But at present only very limited support for highly fragmented files, i.e. ones which have their data attribute split across multiple extents, is included. Another limitation is that at present truncate(2) will never create sparse files, since to mark a file sparse we need to modify the directory entry for the file and we do not implement directory modifications yet.
- The new driver supports write(2) which can both overwrite existing data and extend the file size so that you can write beyond the existing data. Also, writing into sparse regions is supported and the holes are filled in with clusters. But at present only limited support for highly fragmented files, i.e. ones which have their data attribute split across multiple extents, is included. Another limitation is that write(2) will never create sparse files, since to mark a file sparse we need to modify the directory entry for the file and we do not implement directory modifications yet.

### 3.36.4 Supported mount options

In addition to the generic mount options described by the manual page for the mount command (man 8 mount, also see man 5 fstab), the NTFS driver supports the following mount options:

| | |
|---|---|
| iocharset=name | Deprecated option. Still supported but please use nls=name in the future. See description for nls=name. |
| nls=name | Character set to use when returning file names. Unlike VFAT, NTFS suppresses names that contain unconvertible characters. Note that most character sets contain insufficient characters to represent all possible Unicode characters that can exist on NTFS. To be sure you are not missing any files, you are advised to use nls=utf8 which is capable of representing all Unicode characters. |
| utf8=<bool> | Option no longer supported. Currently mapped to nls=utf8 but please use nls=utf8 in the future and make sure utf8 is compiled either as module or into the kernel. See description for nls=name. |
| uid= | |
| gid= | |
| umask= | Provide default owner, group, and access mode mask. These options work as documented in mount(8). By default, the files/directories are owned by root and he/she has read and write permissions, as well as browse permission for directories. No one else has any access permissions. I.e. the mode on all files is by default rw------- and for directories rwx------, a consequence of the default fmask=0177 and dmask=0077. Using a umask of zero will grant all permissions to everyone, i.e. all files and directories will have mode rwxrwxrwx. |
| fmask= | |
| dmask= | Instead of specifying umask which applies both to files and directories, fmask applies only to files and dmask only to directories. |
| sloppy=<BOOL> | If sloppy is specified, ignore unknown mount options. Otherwise the default behaviour is to abort mount if any unknown options are found. |
| show_sys_files=<BOOL> | If show_sys_files is specified, show the system files in directory listings. Otherwise the default behaviour is to hide the system files. Note that even when show_sys_files is specified, "$MFT" will not be visible due to bugs/mis-features in glibc. Further, note that irrespective of show_sys_files, all files are accessible by name, i.e. you can always do "ls -l $UpCase" for example to specifically show the system file containing the Unicode upcase table. |
| case_sensitive=<BOOL> | If case_sensitive is specified, treat all file names as case sensitive and create file names in the POSIX namespace. Otherwise the default behaviour is to treat file names as case insensitive and to create file names in |

### 3.36.5 Known bugs and (mis-)features

- The link count on each directory inode entry is set to 1, due to Linux not supporting directory hard links. This may well confuse some user space applications, since the directory names will have the same inode numbers. This also speeds up ntfs_read_inode() immensely. And we haven't found any problems with this approach so far. If you find a problem with this, please let us know.

Please send bug reports/comments/feedback/abuse to the Linux-NTFS development list at sourceforge: linux-ntfs-dev@lists.sourceforge.net

### 3.36.6 Using NTFS volume and stripe sets

For support of volume and stripe sets, you can either use the kernel's Device-Mapper driver or the kernel's Software RAID / MD driver. The former is the recommended one to use for linear raid. But the latter is required for raid level 5. For striping and mirroring, either driver should work fine.

#### The Device-Mapper driver

You will need to create a table of the components of the volume/stripe set and how they fit together and load this into the kernel using the dmsetup utility (see man 8 dmsetup).

Linear volume sets, i.e. linear raid, has been tested and works fine. Even though untested, there is no reason why stripe sets, i.e. raid level 0, and mirrors, i.e. raid level 1 should not work, too. Stripes with parity, i.e. raid level 5, unfortunately cannot work yet because the current version of the Device-Mapper driver does not support raid level 5. You may be able to use the Software RAID / MD driver for raid level 5, see the next section for details.

To create the table describing your volume you will need to know each of its components and their sizes in sectors, i.e. multiples of 512-byte blocks.

For NT4 fault tolerant volumes you can obtain the sizes using fdisk. So for example if one of your partitions is /dev/hda2 you would do:

```
$ fdisk -ul /dev/hda

Disk /dev/hda: 81.9 GB, 81964302336 bytes
255 heads, 63 sectors/track, 9964 cylinders, total 160086528 sectors
Units = sectors of 1 * 512 = 512 bytes

    Device Boot      Start         End      Blocks   Id  System
    /dev/hda1   *          63     4209029     2104483+  83  Linux
    /dev/hda2         4209030    37768814    16779892+  86  NTFS
    /dev/hda3        37768815    46170809     4200997+  83  Linux
```

And you would know that /dev/hda2 has a size of 37768814 - 4209030 + 1 = 33559785 sectors.

For Win2k and later dynamic disks, you can for example use the ldminfo utility which is part of the Linux LDM tools (the latest version at the time of writing is linux-ldm-0.0.8.tar.bz2). You can download it from:

http://www.linux-ntfs.org/

Simply extract the downloaded archive (tar xvjf linux-ldm-0.0.8.tar.bz2), go into it (cd linux-ldm-0.0.8) and change to the test directory (cd test). You will find the precompiled (i386) ldminfo utility there. NOTE: You will not be able to compile this yourself easily so use the binary version!

Then you would use ldminfo in dump mode to obtain the necessary information:

```
$ ./ldminfo --dump /dev/hda
```

This would dump the LDM database found on /dev/hda which describes all of your dynamic disks and all the volumes on them. At the bottom you will see the VOLUME DEFINITIONS section which is all you really need. You may need to look further above to determine which of the disks in the volume definitions is which device in Linux. Hint: Run ldminfo on each of your dynamic disks and look at the Disk Id close to the top of the output for each (the PRIVATE HEADER section). You can then find these Disk Ids in the VBLK DATABASE section in the <Disk> components where you will get the LDM Name for the disk that is found in the VOLUME DEFINITIONS section.

Note you will also need to enable the LDM driver in the Linux kernel. If your distribution did not enable it, you will need to recompile the kernel with it enabled. This will create the LDM partitions on each device at boot time. You would then use those devices (for /dev/hda they would be /dev/hda1, 2, 3, etc) in the Device-Mapper table.

You can also bypass using the LDM driver by using the main device (e.g. /dev/hda) and then using the offsets of the LDM partitions into this device as the "Start sector of device" when creating the table. Once again ldminfo would give you the correct information to do this.

Assuming you know all your devices and their sizes things are easy.

For a linear raid the table would look like this (note all values are in 512-byte sectors):

```
# Offset into       Size of this    Raid type       Device          Start␣
↪sector
# volume    device                                          of device
0           1028161         linear          /dev/hda1       0
1028161             3903762         linear          /dev/hdb2       0
4931923             2103211         linear          /dev/hdc1       0
```

For a striped volume, i.e. raid level 0, you will need to know the chunk size you used when creating the volume. Windows uses 64kiB as the default, so it will probably be this unless you changes the defaults when creating the array.

For a raid level 0 the table would look like this (note all values are in 512-byte sectors):

```
# Offset   Size         Raid     Number   Chunk  1st        Start   2nd       ␣
↪Start
# into     of the   type     of           size   Device     in      Device    ␣
↪in
# volume   volume        stripes                   device           device
0      2056320  striped  2        128    /dev/hda1  0       /dev/hdb1 0
```

If there are more than two devices, just add each of them to the end of the line.

Finally, for a mirrored volume, i.e. raid level 1, the table would look like this (note all values are in 512-byte sectors):

```
# Ofs Size   Raid   Log  Number Region Should Number Source  Start Target Start
# in  of the type   type of log size   sync?  of     Device  in    Device in
# vol volume                 params             mirrors         Device      ␣
 ↪Device
0    2056320 mirror core 2  16     nosync 2    /dev/hda1 0   /dev/hdb1 0
```

If you are mirroring to multiple devices you can specify further targets at the end of the line.

Note the "Should sync?" parameter "nosync" means that the two mirrors are already in sync which will be the case on a clean shutdown of Windows. If the mirrors are not clean, you can specify the "sync" option instead of "nosync" and the Device-Mapper driver will then copy the entirety of the "Source Device" to the "Target Device" or if you specified multiple target devices to all of them.

Once you have your table, save it in a file somewhere (e.g. /etc/ntfsvolume1), and hand it over to dmsetup to work with, like so:

```
$ dmsetup create myvolume1 /etc/ntfsvolume1
```

You can obviously replace "myvolume1" with whatever name you like.

If it all worked, you will now have the device /dev/device-mapper/myvolume1 which you can then just use as an argument to the mount command as usual to mount the ntfs volume. For example:

```
$ mount -t ntfs -o ro /dev/device-mapper/myvolume1 /mnt/myvol1
```

(You need to create the directory /mnt/myvol1 first and of course you can use anything you like instead of /mnt/myvol1 as long as it is an existing directory.)

It is advisable to do the mount read-only to see if the volume has been setup correctly to avoid the possibility of causing damage to the data on the ntfs volume.

### The Software RAID / MD driver

An alternative to using the Device-Mapper driver is to use the kernel's Software RAID / MD driver. For which you need to set up your /etc/raidtab appropriately (see man 5 raidtab).

Linear volume sets, i.e. linear raid, as well as stripe sets, i.e. raid level 0, have been tested and work fine (though see section "Limitations when using the MD driver with NTFS volumes" especially if you want to use linear raid). Even though untested, there is no reason why mirrors, i.e. raid level 1, and stripes with parity, i.e. raid level 5, should not work, too.

You have to use the "persistent-superblock 0" option for each raid-disk in the NTFS volume/stripe you are configuring in /etc/raidtab as the persistent superblock used by the MD driver would damage the NTFS volume.

Windows by default uses a stripe chunk size of 64k, so you probably want the "chunk-size 64k" option for each raid-disk, too.

For example, if you have a stripe set consisting of two partitions /dev/hda5 and /dev/hdb1 your /etc/raidtab would look like this:

```
raiddev /dev/md0
        raid-level  0
        nr-raid-disks       2
        nr-spare-disks      0
        persistent-superblock       0
        chunk-size  64k
        device              /dev/hda5
        raid-disk   0
        device              /dev/hdb1
        raid-disk   1
```

For linear raid, just change the raid-level above to ”raid-level linear”, for mirrors, change it to ”raid-level 1”, and for stripe sets with parity, change it to ”raid-level 5”.

Note for stripe sets with parity you will also need to tell the MD driver which parity algorithm to use by specifying the option ”parity-algorithm which”, where you need to replace ”which” with the name of the algorithm to use (see man 5 raidtab for available algorithms) and you will have to try the different available algorithms until you find one that works. Make sure you are working read-only when playing with this as you may damage your data otherwise. If you find which algorithm works please let us know (email the linux-ntfs developers list linux-ntfs-dev@lists.sourceforge.net or drop in on IRC in channel #ntfs on the irc.freenode.net network) so we can update this documentation.

Once the raidtab is setup, run for example raid0run -a to start all devices or raid0run /dev/md0 to start a particular md device, in this case /dev/md0.

Then just use the mount command as usual to mount the ntfs volume using for example:

```
mount -t ntfs -o ro /dev/md0 /mnt/myntfsvolume
```

It is advisable to do the mount read-only to see if the md volume has been setup correctly to avoid the possibility of causing damage to the data on the ntfs volume.

### Limitations when using the Software RAID / MD driver

Using the md driver will not work properly if any of your NTFS partitions have an odd number of sectors. This is especially important for linear raid as all data after the first partition with an odd number of sectors will be offset by one or more sectors so if you mount such a partition with write support you will cause massive damage to the data on the volume which will only become apparent when you try to use the volume again under Windows.

So when using linear raid, make sure that all your partitions have an even number of sectors BEFORE attempting to use it. You have been warned!

Even better is to simply use the Device-Mapper for linear raid and then you do not have this problem with odd numbers of sectors.

## 3.37 NTFS3

### 3.37.1 Summary and Features

NTFS3 is fully functional NTFS Read-Write driver. The driver works with NTFS versions up to 3.1. File system type to use on mount is *ntfs3*.

- This driver implements NTFS read/write support for normal, sparse and compressed files.
- Supports native journal replaying.
- Supports NFS export of mounted NTFS volumes.
- Supports extended attributes. Predefined extended attributes:
  - *system.ntfs_security* gets/sets security

    Descriptor: SECURITY_DESCRIPTOR_RELATIVE
  - *system.ntfs_attrib* gets/sets ntfs file/dir attributes.

    Note: Applied to empty files, this allows to switch type between sparse(0x200), compressed(0x800) and normal.
  - *system.ntfs_attrib_be* gets/sets ntfs file/dir attributes.

    Same value as system.ntfs_attrib but always represent as big-endian (endianness of system.ntfs_attrib is the same as of the CPU).

### 3.37.2 Mount Options

The list below describes mount options supported by NTFS3 driver in addition to generic ones. You can use every mount option with **no** option. If it is in this table marked with no it means default is without **no**.

| | |
|---|---|
| iocharset=name | This option informs the driver how to interpret path strings and translate them to Unicode and back. If this option is not set, the default codepage will be used (CONFIG_NLS_DEFAULT).<br>Example: iocharset=utf8 |
| uid= | |
| gid= | |
| umask= | Controls the default permissions for files/directories created after the NTFS volume is mounted. |
| dmask= | Instead of specifying umask which applies both to files and directories, fmask applies only to files and dmask only to directories. |
| fmask= | |
| nohidden | Files with the Windows-specific HIDDEN (FILE_ATTRIBUTE_HIDDEN) attribute will not be shown under Linux. |
| sys_immutable | Files with the Windows-specific SYSTEM (FILE_ATTRIBUTE_SYSTEM) attribute will be marked as system immutable files. |
| hide_dot_files | Updates the Windows-specific HIDDEN (FILE_ATTRIBUTE_HIDDEN) attribute when creating and moving or renaming files. Files whose names start with a dot will have the HIDDEN attribute set and files whose names do not start with a dot will have it unset. |
| windows_names | Prevents the creation of files and directories with a name not allowed by Windows, either because it contains some not allowed character (which are the characters " * / : < > ? \ \| and those whose code is less than 0x20), because the name (with or without extension) is a reserved file name (CON, AUX, NUL, PRN, LPT1-9, COM1-9) or because the last character is a space or a dot. Existing such files can still be read and renamed. |
| discard | Enable support of the TRIM command for improved performance on delete operations, which is recommended for use with the solid-state drives (SSD). |
| force | Forces the driver to mount partitions even if volume is marked dirty. Not recommended for use. |
| sparse | Create new files as sparse. |
| showmeta | Use this parameter to show all meta-files (System Files) on a mounted NTFS partition. By default, all meta-files are hidden. |
| prealloc | Preallocate space for files excessively when file size is increasing on writes. Decreases fragmentation in case of parallel write operations to different files. |
| acl | Support POSIX ACLs (Access Control Lists). Effective if supported by Kernel. Not to be confused with NTFS ACLs. The option specified as acl enables support for POSIX ACLs. |

### 3.37.3 Todo list

- Full journaling support over JBD. Currently journal replaying is supported which is not necessarily as effective as JBD would be.

### 3.37.4 References

- **Commercial version of the NTFS driver for Linux.**
  https://www.paragon-software.com/home/ntfs-linux-professional/
- **Direct e-mail address for feedback and requests on the NTFS3 implementation.**
  almaz.alexandrovich@paragon-software.com

## 3.38 OCFS2 filesystem

OCFS2 is a general purpose extent based shared disk cluster file system with many similarities to ext3. It supports 64 bit inode numbers, and has automatically extending metadata groups which may also make it attractive for non-clustered use.

You'll want to install the ocfs2-tools package in order to at least get ”mount.ocfs2” and ”ocfs2_hb_ctl”.

Project web page: http://ocfs2.wiki.kernel.org Tools git tree: https://github.com/markfasheh/ocfs2-tools OCFS2 mailing lists: https://subspace.kernel.org/lists.linux.dev.html



### 3.38.1 Credits

Lots of code taken from ext3 and other projects.

Authors in alphabetical order:

- Joel Becker <joel.becker@oracle.com>
- Zach Brown <zach.brown@oracle.com>
- Mark Fasheh <mfasheh@suse.com>
- Kurt Hackel <kurt.hackel@oracle.com>
- Tao Ma <tao.ma@oracle.com>
- Sunil Mushran <sunil.mushran@oracle.com>
- Manish Singh <manish.singh@oracle.com>
- Tiger Yang <tiger.yang@oracle.com>

### 3.38.2 Caveats

Features which OCFS2 does not support yet:

- Directory change notification (F_NOTIFY)
- Distributed Caching (F_SETLEASE/F_GETLEASE/break_lease)

### 3.38.3 Mount options

OCFS2 supports the following mount options:

(*) == default

| | |
|---|---|
| barrier=1 | This enables/disables barriers. barrier=0 disables it, barrier=1 enables it. |
| errors=remount-ro(*) | Remount the filesystem read-only on an error. |
| errors=panic | Panic and halt the machine if an error occurs. |
| intr (*) | Allow signals to interrupt cluster operations. |
| nointr | Do not allow signals to interrupt cluster operations. |
| noatime | Do not update access time. |
| relatime(*) | Update atime if the previous atime is older than mtime or ctime |
| strictatime | Always update atime, but the minimum update interval is specified by atime_quantum. |
| atime_quantum=60(*) | OCFS2 will not update atime unless this number of seconds has passed since the last update. Set to zero to always update atime. This option need work with strictatime. |
| data=ordered (*) | All data are forced directly out to the main file system prior to its metadata being committed to the journal. |
| data=writeback | Data ordering is not preserved, data may be written into the main file system after its metadata has been committed to the journal. |
| preferred_slot=0(*) | During mount, try to use this filesystem slot first. If it is in use by another node, the first empty one found will be chosen. Invalid values will be ignored. |
| commit=nrsec (*) | Ocfs2 can be told to sync all its data and metadata every 'nrsec' seconds. The default value is 5 seconds. This means that if you lose your power, you will lose as much as the latest 5 seconds of work (your filesystem will not be damaged though, thanks to the journaling). This default value (or any low value) will hurt performance, but it's good for data-safety. Setting it to 0 will have the same effect as leaving it at the default (5 seconds). Setting it to very large values will improve performance. |
| localalloc=8(*) | Allows custom localalloc size in MB. If the value is too large, the fs will silently revert it to the default. |
| localflocks | This disables cluster aware flock. |
| inode64 | Indicates that Ocfs2 is allowed to create inodes at any location in the filesystem, including those which will result in inode numbers occupying more than 32 bits of significance. |
| user_xattr (*) | Enables Extended User Attributes. |
| nouser_xattr | Disables Extended User Attributes. |
| acl | Enables POSIX Access Control Lists support. |
| noacl (*) | Disables POSIX Access Control Lists support. |
| resv_level=2 (*) | Set how aggressive allocation reservations will be. Valid values are between 0 (reservations off) to 8 (maximum space for reservations). |
| dir_resv_level= (*) | By default, directory reservations will scale with file reservations - users should rarely need to change this value. If allocation reservations are turned off, this option will have no effect. |
| coherency=full (*) | Disallow concurrent O_DIRECT writes, cluster inode lock will be taken to force other nodes drop cache, therefore full cluster coherency is guaranteed even for O_DIRECT writes. |
| coherency=buffered | Allow concurrent O_DIRECT writes without EX lock among nodes, which gains high performance at risk of getting stale data on other nodes. |
| journal_async_commit | Commit block can be written to disk without waiting for descriptor blocks. If enabled older kernels cannot mount the device. This will enable 'journal_checksum' internally. |

## 3.39 OCFS2 file system - online file check

This document will describe OCFS2 online file check feature.

### 3.39.1 Introduction

OCFS2 is often used in high-availability systems. However, OCFS2 usually converts the filesystem to read-only when encounters an error. This may not be necessary, since turning the filesystem read-only would affect other running processes as well, decreasing availability. Then, a mount option (errors=continue) is introduced, which would return the -EIO errno to the calling process and terminate further processing so that the filesystem is not corrupted further. The filesystem is not converted to read-only, and the problematic file's inode number is reported in the kernel log. The user can try to check/fix this file via online filecheck feature.

### 3.39.2 Scope

This effort is to check/fix small issues which may hinder day-to-day operations of a cluster filesystem by turning the filesystem read-only. The scope of checking/fixing is at the file level, initially for regular files and eventually to all files (including system files) of the filesystem.

In case of directory to file links is incorrect, the directory inode is reported as erroneous.

This feature is not suited for extravagant checks which involve dependency of other components of the filesystem, such as but not limited to, checking if the bits for file blocks in the allocation has been set. In case of such an error, the offline fsck should/would be recommended.

Finally, such an operation/feature should not be automated lest the filesystem may end up with more damage than before the repair attempt. So, this has to be performed using user interaction and consent.

### 3.39.3 User interface

When there are errors in the OCFS2 filesystem, they are usually accompanied by the inode number which caused the error. This inode number would be the input to check/fix the file.

There is a sysfs directory for each OCFS2 file system mounting:

```
/sys/fs/ocfs2/<devname>/filecheck
```

Here, <devname> indicates the name of OCFS2 volume device which has been already mounted. The file above would accept inode numbers. This could be used to communicate with kernel space, tell which file(inode number) will be checked or fixed. Currently, three operations are supported, which includes checking inode, fixing inode and setting the size of result record history.

1. If you want to know what error exactly happened to <inode> before fixing, do:

```
# echo "<inode>" > /sys/fs/ocfs2/<devname>/filecheck/check
# cat /sys/fs/ocfs2/<devname>/filecheck/check
```

The output is like this:

```
INO         DONE    ERROR
39502               1       GENERATION

<INO> lists the inode numbers.
<DONE> indicates whether the operation has been finished.
<ERROR> says what kind of errors was found. For the detailed error numbers,
please refer to the file linux/fs/ocfs2/filecheck.h.
```

2. If you determine to fix this inode, do:

```
# echo "<inode>" > /sys/fs/ocfs2/<devname>/filecheck/fix
# cat /sys/fs/ocfs2/<devname>/filecheck/fix
```

The output is like this::

```
INO         DONE    ERROR
39502               1       SUCCESS
```

This time, the <ERROR> column indicates whether this fix is successful or not.

3. The record cache is used to store the history of check/fix results. It's default size is 10, and can be adjust between the range of 10 ~ 100. You can adjust the size like this:

```
# echo "<size>" > /sys/fs/ocfs2/<devname>/filecheck/set
```

### 3.39.4 Fixing stuff

On receiving the inode, the filesystem would read the inode and the file metadata. In case of errors, the filesystem would fix the errors and report the problems it fixed in the kernel log. As a precautionary measure, the inode must first be checked for errors before performing a final fix.

The inode and the result history will be maintained temporarily in a small linked list buffer which would contain the last (N) inodes fixed/checked, the detailed errors which were fixed/checked are printed in the kernel log.

## 3.40 Optimized MPEG Filesystem (OMFS)

### 3.40.1 Overview

OMFS is a filesystem created by SonicBlue for use in the ReplayTV DVR and Rio Karma MP3 player. The filesystem is extent-based, utilizing block sizes from 2k to 8k, with hash-based directories. This filesystem driver may be used to read and write disks from these devices.

Note, it is not recommended that this FS be used in place of a general filesystem for your own streaming media device. Native Linux filesystems will likely perform better.

More information is available at:

http://linux-karma.sf.net/

Various utilities, including mkomfs and omfsck, are included with omfsprogs, available at:

https://bobcopeland.com/karma/

Instructions are included in its README.

### 3.40.2 Options

OMFS supports the following mount-time options:

| | |
|---|---|
| uid=n | make all files owned by specified user |
| gid=n | make all files owned by specified group |
| umask=xxx | set permission umask to xxx |
| fmask=xxx | set umask to xxx for files |
| dmask=xxx | set umask to xxx for directories |

### 3.40.3 Disk format

OMFS discriminates between ”sysblocks” and normal data blocks. The sysblock group consists of super block information, file metadata, directory structures, and extents. Each sysblock has a header containing CRCs of the entire sysblock, and may be mirrored in successive blocks on the disk. A sysblock may have a smaller size than a data block, but since they are both addressed by the same 64-bit block number, any remaining space in the smaller sysblock is unused.

Sysblock header information:

```
struct omfs_header {
        __be64 h_self;                  /* FS block where this is located */
        __be32 h_body_size;             /* size of useful data after header */
        __be16 h_crc;                   /* crc-ccitt of body_size bytes */
        char h_fill1[2];
        u8 h_version;                   /* version, always 1 */
        char h_type;                    /* OMFS_INODE_X */
        u8 h_magic;                     /* OMFS_IMAGIC */
        u8 h_check_xor;                 /* XOR of header bytes before this */
        __be32 h_fill2;
};
```

Files and directories are both represented by omfs_inode:

```
struct omfs_inode {
        struct omfs_header i_head;      /* header */
        __be64 i_parent;                /* parent containing this inode */
        __be64 i_sibling;               /* next inode in hash bucket */
        __be64 i_ctime;                 /* ctime, in milliseconds */
        char i_fill1[35];
        char i_type;                    /* OMFS_[DIR,FILE] */
        __be32 i_fill2;
        char i_fill3[64];
        char i_name[OMFS_NAMELEN];      /* filename */
        __be64 i_size;                  /* size of file, in bytes */
};
```

Directories in OMFS are implemented as a large hash table. Filenames are hashed then prepended into the bucket list beginning at OMFS_DIR_START. Lookup requires hashing the filename, then seeking across i_sibling pointers until a match is found on i_name. Empty buckets are represented by block pointers with all-1s (~0).

A file is an omfs_inode structure followed by an extent table beginning at OMFS_EXTENT_START:

```
struct omfs_extent_entry {
        __be64 e_cluster;               /* start location of a set of blocks */
        __be64 e_blocks;                /* number of blocks after e_cluster */
};

struct omfs_extent {
        __be64 e_next;                  /* next extent table location */
        __be32 e_extent_count;          /* total # extents in this table */
        __be32 e_fill;
        struct omfs_extent_entry e_entry;       /* start of extent entries */
};
```

Each extent holds the block offset followed by number of blocks allocated to the extent. The final extent in each table is a terminator with e_cluster being ~0 and e_blocks being ones'-complement of the total number of blocks in the table.

If this table overflows, a continuation inode is written and pointed to by e_next. These have a header but lack the rest of the inode structure.

## 3.41 ORANGEFS

OrangeFS is an LGPL userspace scale-out parallel storage system. It is ideal for large storage problems faced by HPC, BigData, Streaming Video, Genomics, Bioinformatics.

Orangefs, originally called PVFS, was first developed in 1993 by Walt Ligon and Eric Blumer as a parallel file system for Parallel Virtual Machine (PVM) as part of a NASA grant to study the I/O patterns of parallel programs.

Orangefs features include:

- Distributes file data among multiple file servers
- Supports simultaneous access by multiple clients
- Stores file data and metadata on servers using local file system and access methods
- Userspace implementation is easy to install and maintain
- Direct MPI support
- Stateless

### 3.41.1 Mailing List Archives

http://lists.orangefs.org/pipermail/devel_lists.orangefs.org/

### 3.41.2 Mailing List Submissions

devel@lists.orangefs.org

### 3.41.3 Documentation

http://www.orangefs.org/documentation/

### 3.41.4 Running ORANGEFS On a Single Server

OrangeFS is usually run in large installations with multiple servers and clients, but a complete filesystem can be run on a single machine for development and testing.

On Fedora, install orangefs and orangefs-server:

```
dnf -y install orangefs orangefs-server
```

There is an example server configuration file in /etc/orangefs/orangefs.conf. Change localhost to your hostname if necessary.

To generate a filesystem to run xfstests against, see below.

There is an example client configuration file in /etc/pvfs2tab. It is a single line. Uncomment it and change the hostname if necessary. This controls clients which use libpvfs2. This does not control the pvfs2-client-core.

Create the filesystem:

```
pvfs2-server -f /etc/orangefs/orangefs.conf
```

Start the server:

```
systemctl start orangefs-server
```

Test the server:

```
pvfs2-ping -m /pvfsmnt
```

Start the client. The module must be compiled in or loaded before this point:

```
systemctl start orangefs-client
```

Mount the filesystem:

```
mount -t pvfs2 tcp://localhost:3334/orangefs /pvfsmnt
```

### 3.41.5 Userspace Filesystem Source

http://www.orangefs.org/download

Orangefs versions prior to 2.9.3 would not be compatible with the upstream version of the kernel client.

### 3.41.6 Building ORANGEFS on a Single Server

Where OrangeFS cannot be installed from distribution packages, it may be built from source.

You can omit --prefix if you don't care that things are sprinkled around in /usr/local. As of version 2.9.6, OrangeFS uses Berkeley DB by default, we will probably be changing the default to LMDB soon.

```
./configure --prefix=/opt/ofs --with-db-backend=lmdb --disable-usrint

make

make install
```

Create an orangefs config file by running pvfs2-genconfig and specifying a target config file. Pvfs2-genconfig will prompt you through. Generally it works fine to take the defaults, but you should use your server's hostname, rather than ”localhost” when it comes to that question:

```
/opt/ofs/bin/pvfs2-genconfig /etc/pvfs2.conf
```

Create an /etc/pvfs2tab file (localhost is fine):

```
echo tcp://localhost:3334/orangefs /pvfsmnt pvfs2 defaults,noauto 0 0 > \
    /etc/pvfs2tab
```

Create the mount point you specified in the tab file if needed:

```
mkdir /pvfsmnt
```

Bootstrap the server:

```
/opt/ofs/sbin/pvfs2-server -f /etc/pvfs2.conf
```

Start the server:

```
/opt/ofs/sbin/pvfs2-server /etc/pvfs2.conf
```

Now the server should be running. Pvfs2-ls is a simple test to verify that the server is running:

```
/opt/ofs/bin/pvfs2-ls /pvfsmnt
```

If stuff seems to be working, load the kernel module and turn on the client core:

```
/opt/ofs/sbin/pvfs2-client -p /opt/ofs/sbin/pvfs2-client-core
```

Mount your filesystem:

```
mount -t pvfs2 tcp://`hostname`:3334/orangefs /pvfsmnt
```

### 3.41.7 Running xfstests

It is useful to use a scratch filesystem with xfstests. This can be done with only one server.

Make a second copy of the FileSystem section in the server configuration file, which is /etc/orangefs/orangefs.conf. Change the Name to scratch. Change the ID to something other than the ID of the first FileSystem section (2 is usually a good choice).

Then there are two FileSystem sections: orangefs and scratch.

This change should be made before creating the filesystem.

```
pvfs2-server -f /etc/orangefs/orangefs.conf
```

To run xfstests, create /etc/xfsqa.config:

```
TEST_DIR=/orangefs
TEST_DEV=tcp://localhost:3334/orangefs
SCRATCH_MNT=/scratch
SCRATCH_DEV=tcp://localhost:3334/scratch
```

Then xfstests can be run:

```
./check -pvfs2
```

### 3.41.8 Options

The following mount options are accepted:

**acl**
: Allow the use of Access Control Lists on files and directories.

**intr**
: Some operations between the kernel client and the user space filesystem can be interruptible, such as changes in debug levels and the setting of tunable parameters.

**local_lock**
: Enable posix locking from the perspective of ”this” kernel. The default file_operations lock action is to return ENOSYS. Posix locking kicks in if the filesystem is mounted with -o local_lock. Distributed locking is being worked on for the future.

### 3.41.9 Debugging

If you want the debug (GOSSIP) statements in a particular source file (inode.c for example) go to syslog:

```
echo inode > /sys/kernel/debug/orangefs/kernel-debug
```

No debugging (the default):

```
echo none > /sys/kernel/debug/orangefs/kernel-debug
```

Debugging from several source files:

```
echo inode,dir > /sys/kernel/debug/orangefs/kernel-debug
```

All debugging:

```
echo all > /sys/kernel/debug/orangefs/kernel-debug
```

Get a list of all debugging keywords:

```
cat /sys/kernel/debug/orangefs/debug-help
```

### 3.41.10 Protocol between Kernel Module and Userspace

Orangefs is a user space filesystem and an associated kernel module. We'll just refer to the user space part of Orangefs as ”userspace” from here on out. Orangefs descends from PVFS, and userspace code still uses PVFS for function and variable names. Userspace typedefs many of the important structures. Function and variable names in the kernel module have been transitioned to ”orangefs”, and The Linux Coding Style avoids typedefs, so kernel module structures that correspond to userspace structures are not typedefed.

The kernel module implements a pseudo device that userspace can read from and write to. Userspace can also manipulate the kernel module through the pseudo device with ioctl.

#### The Bufmap

At startup userspace allocates two page-size-aligned (posix_memalign) mlocked memory buffers, one is used for IO and one is used for readdir operations. The IO buffer is 41943040 bytes and the readdir buffer is 4194304 bytes. Each buffer contains logical chunks, or partitions, and a pointer to each buffer is added to its own PVFS_dev_map_desc structure which also describes its total size, as well as the size and number of the partitions.

A pointer to the IO buffer's PVFS_dev_map_desc structure is sent to a mapping routine in the kernel module with an ioctl. The structure is copied from user space to kernel space with copy_from_user and is used to initialize the kernel module's ”bufmap” (struct orangefs_bufmap), which then contains:

- refcnt - a reference counter
- desc_size - PVFS2_BUFMAP_DEFAULT_DESC_SIZE (4194304) - the IO buffer's partition size, which represents the filesystem's block size and is used for s_blocksize in super blocks.

- desc_count - PVFS2_BUFMAP_DEFAULT_DESC_COUNT (10) - the number of partitions in the IO buffer.
- desc_shift - log2(desc_size), used for s_blocksize_bits in super blocks.
- total_size - the total size of the IO buffer.
- page_count - the number of 4096 byte pages in the IO buffer.
- page_array - a pointer to `page_count * (sizeof(struct page*))` bytes of kcalloced memory. This memory is used as an array of pointers to each of the pages in the IO buffer through a call to get_user_pages.
- desc_array - a pointer to `desc_count * (sizeof(struct orangefs_bufmap_desc))` bytes of kcalloced memory. This memory is further initialized:

  user_desc is the kernel's copy of the IO buffer's ORANGEFS_dev_map_desc structure. user_desc->ptr points to the IO buffer.

```
pages_per_desc = bufmap->desc_size / PAGE_SIZE
offset = 0

bufmap->desc_array[0].page_array = &bufmap->page_array[offset]
bufmap->desc_array[0].array_count = pages_per_desc = 1024
bufmap->desc_array[0].uaddr = (user_desc->ptr) + (0 * 1024 * 4096)
offset += 1024
                    .
                    .
                    .
bufmap->desc_array[9].page_array = &bufmap->page_array[offset]
bufmap->desc_array[9].array_count = pages_per_desc = 1024
bufmap->desc_array[9].uaddr = (user_desc->ptr) +
                                        (9 * 1024 * 4096)
offset += 1024
```

- buffer_index_array - a desc_count sized array of ints, used to indicate which of the IO buffer's partitions are available to use.
- buffer_index_lock - a spinlock to protect buffer_index_array during update.
- readdir_index_array - a five (ORANGEFS_READDIR_DEFAULT_DESC_COUNT) element int array used to indicate which of the readdir buffer's partitions are available to use.
- readdir_index_lock - a spinlock to protect readdir_index_array during update.

### Operations

The kernel module builds an ”op” (struct orangefs_kernel_op_s) when it needs to communicate with userspace. Part of the op contains the ”upcall” which expresses the request to userspace. Part of the op eventually contains the ”downcall” which expresses the results of the request.

The slab allocator is used to keep a cache of op structures handy.

At init time the kernel module defines and initializes a request list and an in_progress hash table to keep track of all the ops that are in flight at any given time.

Ops are stateful:

- **unknown**
  - op was just initialized
- **waiting**
  - op is on request_list (upward bound)
- **inprogr**
  - op is in progress (waiting for downcall)
- **serviced**
  - op has matching downcall; ok
- **purged**
  - op has to start a timer since client-core exited uncleanly before servicing op
- **given up**
  - submitter has given up waiting for it

When some arbitrary userspace program needs to perform a filesystem operation on Orangefs (readdir, I/O, create, whatever) an op structure is initialized and tagged with a distinguishing ID number. The upcall part of the op is filled out, and the op is passed to the "service_operation" function.

Service_operation changes the op's state to "waiting", puts it on the request list, and signals the Orangefs file_operations.poll function through a wait queue. Userspace is polling the pseudo-device and thus becomes aware of the upcall request that needs to be read.

When the Orangefs file_operations.read function is triggered, the request list is searched for an op that seems ready-to-process. The op is removed from the request list. The tag from the op and the filled-out upcall struct are copy_to_user'ed back to userspace.

If any of these (and some additional protocol) copy_to_users fail, the op's state is set to "waiting" and the op is added back to the request list. Otherwise, the op's state is changed to "in progress", and the op is hashed on its tag and put onto the end of a list in the in_progress hash table at the index the tag hashed to.

When userspace has assembled the response to the upcall, it writes the response, which includes the distinguishing tag, back to the pseudo device in a series of io_vecs. This triggers the Orangefs file_operations.write_iter function to find the op with the associated tag and remove it from the in_progress hash table. As long as the op's state is not "canceled" or "given up", its state is set to "serviced". The file_operations.write_iter function returns to the waiting vfs, and back to service_operation through wait_for_matching_downcall.

Service operation returns to its caller with the op's downcall part (the response to the upcall) filled out.

The "client-core" is the bridge between the kernel module and userspace. The client-core is a daemon. The client-core has an associated watchdog daemon. If the client-core is ever signaled to die, the watchdog daemon restarts the client-core. Even though the client-core is restarted "right away", there is a period of time during such an event that the client-core is dead. A dead client-core can't be triggered by the Orangefs file_operations.poll function. Ops that pass through service_operation during a "dead spell" can timeout on the wait queue and one attempt is made to recycle them. Obviously, if the client-core stays dead too long, the arbitrary userspace processes trying to use Orangefs will be negatively affected. Waiting ops that can't

be serviced will be removed from the request list and have their states set to ”given up”. In-progress ops that can't be serviced will be removed from the in_progress hash table and have their states set to ”given up”.

Readdir and I/O ops are atypical with respect to their payloads.

- readdir ops use the smaller of the two pre-allocated pre-partitioned memory buffers. The readdir buffer is only available to userspace. The kernel module obtains an index to a free partition before launching a readdir op. Userspace deposits the results into the indexed partition and then writes them to back to the pvfs device.
- io (read and write) ops use the larger of the two pre-allocated pre-partitioned memory buffers. The IO buffer is accessible from both userspace and the kernel module. The kernel module obtains an index to a free partition before launching an io op. The kernel module deposits write data into the indexed partition, to be consumed directly by userspace. Userspace deposits the results of read requests into the indexed partition, to be consumed directly by the kernel module.

Responses to kernel requests are all packaged in pvfs2_downcall_t structs. Besides a few other members, pvfs2_downcall_t contains a union of structs, each of which is associated with a particular response type.

The several members outside of the union are:

**int32_t type**

- type of operation.

**int32_t status**

- return code for the operation.

**int64_t trailer_size**

- 0 unless readdir operation.

**char *trailer_buf**

- initialized to NULL, used during readdir operations.

The appropriate member inside the union is filled out for any particular response.

**PVFS2_VFS_OP_FILE_IO**
fill a pvfs2_io_response_t

**PVFS2_VFS_OP_LOOKUP**
fill a PVFS_object_kref

**PVFS2_VFS_OP_CREATE**
fill a PVFS_object_kref

**PVFS2_VFS_OP_SYMLINK**
fill a PVFS_object_kref

**PVFS2_VFS_OP_GETATTR**
fill in a PVFS_sys_attr_s (tons of stuff the kernel doesn't need) fill in a string with the link target when the object is a symlink.

**PVFS2_VFS_OP_MKDIR**
fill a PVFS_object_kref

**PVFS2_VFS_OP_STATFS**
fill a pvfs2_statfs_response_t with useless info <g>. It is hard for us to know, in a timely fashion, these statistics about our distributed network filesystem.

**PVFS2_VFS_OP_FS_MOUNT**
fill a pvfs2_fs_mount_response_t which is just like a PVFS_object_kref except its members are in a different order and "__pad1" is replaced with "id".

**PVFS2_VFS_OP_GETXATTR**
fill a pvfs2_getxattr_response_t

**PVFS2_VFS_OP_LISTXATTR**
fill a pvfs2_listxattr_response_t

**PVFS2_VFS_OP_PARAM**
fill a pvfs2_param_response_t

**PVFS2_VFS_OP_PERF_COUNT**
fill a pvfs2_perf_count_response_t

**PVFS2_VFS_OP_FSKEY**
file a pvfs2_fs_key_response_t

**PVFS2_VFS_OP_READDIR**
jamb everything needed to represent a pvfs2_readdir_response_t into the readdir buffer descriptor specified in the upcall.

Userspace uses writev() on /dev/pvfs2-req to pass responses to the requests made by the kernel side.

A buffer_list containing:

- a pointer to the prepared response to the request from the kernel (struct pvfs2_downcall_t).
- and also, in the case of a readdir request, a pointer to a buffer containing descriptors for the objects in the target directory.

... is sent to the function (PINT_dev_write_list) which performs the writev.

PINT_dev_write_list has a local iovec array: struct iovec io_array[10];

The first four elements of io_array are initialized like this for all responses:

```
io_array[0].iov_base = address of local variable "proto_ver" (int32_t)
io_array[0].iov_len = sizeof(int32_t)

io_array[1].iov_base = address of global variable "pdev_magic" (int32_t)
io_array[1].iov_len = sizeof(int32_t)

io_array[2].iov_base = address of parameter "tag" (PVFS_id_gen_t)
io_array[2].iov_len = sizeof(int64_t)

io_array[3].iov_base = address of out_downcall member (pvfs2_downcall_t)
                       of global variable vfs_request (vfs_request_t)
io_array[3].iov_len = sizeof(pvfs2_downcall_t)
```

Readdir responses initialize the fifth element io_array like this:

```
io_array[4].iov_base = contents of member trailer_buf (char *)
                        from out_downcall member of global variable
                        vfs_request
io_array[4].iov_len = contents of member trailer_size (PVFS_size)
                       from out_downcall member of global variable
                       vfs_request
```

Orangefs exploits the dcache in order to avoid sending redundant requests to userspace. We keep object inode attributes up-to-date with orangefs_inode_getattr. Orangefs_inode_getattr uses two arguments to help it decide whether or not to update an inode: ”new” and ”bypass”. Orangefs keeps private data in an object's inode that includes a short timeout value, getattr_time, which allows any iteration of orangefs_inode_getattr to know how long it has been since the inode was updated. When the object is not new (new == 0) and the bypass flag is not set (bypass == 0) orangefs_inode_getattr returns without updating the inode if getattr_time has not timed out. Getattr_time is updated each time the inode is updated.

Creation of a new object (file, dir, sym-link) includes the evaluation of its pathname, resulting in a negative directory entry for the object. A new inode is allocated and associated with the dentry, turning it from a negative dentry into a ”productive full member of society”. Orangefs obtains the new inode from Linux with *new_inode()* and associates the inode with the dentry by sending the pair back to Linux with *d_instantiate()*.

The evaluation of a pathname for an object resolves to its corresponding dentry. If there is no corresponding dentry, one is created for it in the dcache. Whenever a dentry is modified or verified Orangefs stores a short timeout value in the dentry's d_time, and the dentry will be trusted for that amount of time. Orangefs is a network filesystem, and objects can potentially change out-of-band with any particular Orangefs kernel module instance, so trusting a dentry is risky. The alternative to trusting dentries is to always obtain the needed information from userspace - at least a trip to the client-core, maybe to the servers. Obtaining information from a dentry is cheap, obtaining it from userspace is relatively expensive, hence the motivation to use the dentry when possible.

The timeout values d_time and getattr_time are jiffy based, and the code is designed to avoid the jiffy-wrap problem:

```
"In general, if the clock may have wrapped around more than once, there
is no way to tell how much time has elapsed. However, if the times t1
and t2 are known to be fairly close, we can reliably compute the
difference in a way that takes into account the possibility that the
clock may have wrapped between times."
```

from course notes by instructor Andy Wang

Written by: Neil Brown Please see MAINTAINERS file for where to send questions.

## 3.42 Overlay Filesystem

This document describes a prototype for a new approach to providing overlay-filesystem functionality in Linux (sometimes referred to as union-filesystems). An overlay-filesystem tries to present a filesystem which is the result over overlaying one filesystem on top of the other.

### 3.42.1 Overlay objects

The overlay filesystem approach is 'hybrid', because the objects that appear in the filesystem do not always appear to belong to that filesystem. In many cases, an object accessed in the union will be indistinguishable from accessing the corresponding object from the original filesystem. This is most obvious from the 'st_dev' field returned by stat(2).

While directories will report an st_dev from the overlay-filesystem, non-directory objects may report an st_dev from the lower filesystem or upper filesystem that is providing the object. Similarly st_ino will only be unique when combined with st_dev, and both of these can change over the lifetime of a non-directory object. Many applications and tools ignore these values and will not be affected.

In the special case of all overlay layers on the same underlying filesystem, all objects will report an st_dev from the overlay filesystem and st_ino from the underlying filesystem. This will make the overlay mount more compliant with filesystem scanners and overlay objects will be distinguishable from the corresponding objects in the original filesystem.

On 64bit systems, even if all overlay layers are not on the same underlying filesystem, the same compliant behavior could be achieved with the ”xino” feature. The ”xino” feature composes a unique object identifier from the real object st_ino and an underlying fsid number. The ”xino” feature uses the high inode number bits for fsid, because the underlying filesystems rarely use the high inode number bits. In case the underlying inode number does overflow into the high xino bits, overlay filesystem will fall back to the non xino behavior for that inode.

The ”xino” feature can be enabled with the ”-o xino=on” overlay mount option. If all underlying filesystems support NFS file handles, the value of st_ino for overlay filesystem objects is not only unique, but also persistent over the lifetime of the filesystem. The ”-o xino=auto” overlay mount option enables the ”xino” feature only if the persistent st_ino requirement is met.

The following table summarizes what can be expected in different overlay configurations.

#### Inode properties

| Configuration | Persistent st_ino | | Uniform st_dev | | st_ino == d_ino | | d_ino == i_ino [*] | |
|---|---|---|---|---|---|---|---|---|
| | dir | !dir | dir | !dir | dir | !dir | dir | !dir |
| All layers on same fs | Y | Y | Y | Y | Y | Y | Y | Y |
| Layers not on same fs, xino=off | N | N | Y | N | N | Y | N | Y |
| xino=on/auto | Y | Y | Y | Y | Y | Y | Y | Y |
| xino=on/auto, ino overflow | N | N | Y | N | N | Y | N | Y |

[*] nfsd v3 readdirplus verifies d_ino == i_ino. i_ino is exposed via several /proc files, such as /proc/locks and /proc/self/fdinfo/<fd> of an inotify file descriptor.

### 3.42.2 Upper and Lower

An overlay filesystem combines two filesystems - an 'upper' filesystem and a 'lower' filesystem. When a name exists in both filesystems, the object in the 'upper' filesystem is visible while the object in the 'lower' filesystem is either hidden or, in the case of directories, merged with the 'upper' object.

It would be more correct to refer to an upper and lower 'directory tree' rather than 'filesystem' as it is quite possible for both directory trees to be in the same filesystem and there is no requirement that the root of a filesystem be given for either upper or lower.

A wide range of filesystems supported by Linux can be the lower filesystem, but not all filesystems that are mountable by Linux have the features needed for OverlayFS to work. The lower filesystem does not need to be writable. The lower filesystem can even be another overlayfs. The upper filesystem will normally be writable and if it is it must support the creation of trusted.* and/or user.* extended attributes, and must provide valid d_type in readdir responses, so NFS is not suitable.

A read-only overlay of two read-only filesystems may use any filesystem type.

### 3.42.3 Directories

Overlaying mainly involves directories. If a given name appears in both upper and lower filesystems and refers to a non-directory in either, then the lower object is hidden - the name refers only to the upper object.

Where both upper and lower objects are directories, a merged directory is formed.

At mount time, the two directories given as mount options ”lowerdir” and ”upperdir” are combined into a merged directory:

```
mount -t overlay overlay -olowerdir=/lower,upperdir=/upper,\
workdir=/work /merged
```

The ”workdir” needs to be an empty directory on the same filesystem as upperdir.

Then whenever a lookup is requested in such a merged directory, the lookup is performed in each actual directory and the combined result is cached in the dentry belonging to the overlay filesystem. If both actual lookups find directories, both are stored and a merged directory is created, otherwise only one is stored: the upper if it exists, else the lower.

Only the lists of names from directories are merged. Other content such as metadata and extended attributes are reported for the upper directory only. These attributes of the lower directory are hidden.

### 3.42.4 whiteouts and opaque directories

In order to support rm and rmdir without changing the lower filesystem, an overlay filesystem needs to record in the upper filesystem that files have been removed. This is done using whiteouts and opaque directories (non-directories are always opaque).

A whiteout is created as a character device with 0/0 device number or as a zero-size regular file with the xattr ”trusted.overlay.whiteout”.

When a whiteout is found in the upper level of a merged directory, any matching name in the lower level is ignored, and the whiteout itself is also hidden.

A directory is made opaque by setting the xattr ”trusted.overlay.opaque” to ”y”. Where the upper filesystem contains an opaque directory, any directory in the lower filesystem with the same name is ignored.

An opaque directory should not conntain any whiteouts, because they do not serve any purpose. A merge directory containing regular files with the xattr ”trusted.overlay.whiteout”, should be additionally marked by setting the xattr ”trusted.overlay.opaque” to ”x” on the merge directory itself. This is needed to avoid the overhead of checking the ”trusted.overlay.whiteout” on all entries during readdir in the common case.

### 3.42.5 readdir

When a 'readdir' request is made on a merged directory, the upper and lower directories are each read and the name lists merged in the obvious way (upper is read first, then lower - entries that already exist are not re-added). This merged name list is cached in the 'struct file' and so remains as long as the file is kept open. If the directory is opened and read by two processes at the same time, they will each have separate caches. A seekdir to the start of the directory (offset 0) followed by a readdir will cause the cache to be discarded and rebuilt.

This means that changes to the merged directory do not appear while a directory is being read. This is unlikely to be noticed by many programs.

seek offsets are assigned sequentially when the directories are read. Thus if:

- read part of a directory
- remember an offset, and close the directory
- re-open the directory some time later

- seek to the remembered offset

there may be little correlation between the old and new locations in the list of filenames, particularly if anything has changed in the directory.

Readdir on directories that are not merged is simply handled by the underlying directory (upper or lower).

### 3.42.6 renaming directories

When renaming a directory that is on the lower layer or merged (i.e. the directory was not created on the upper layer to start with) overlayfs can handle it in two different ways:

1. return EXDEV error: this error is returned by rename(2) when trying to move a file or directory across filesystem boundaries. Hence applications are usually prepared to handle this error (mv(1) for example recursively copies the directory tree). This is the default behavior.
2. If the ”redirect_dir” feature is enabled, then the directory will be copied up (but not the contents). Then the ”trusted.overlay.redirect” extended attribute is set to the path of the original location from the root of the overlay. Finally the directory is moved to the new location.

There are several ways to tune the ”redirect_dir” feature.

Kernel config options:

- **OVERLAY_FS_REDIRECT_DIR:**
  If this is enabled, then redirect_dir is turned on by default.
- **OVERLAY_FS_REDIRECT_ALWAYS_FOLLOW:**
  If this is enabled, then redirects are always followed by default. Enabling this results in a less secure configuration. Enable this option only when worried about backward compatibility with kernels that have the redirect_dir feature and follow redirects even if turned off.

Module options (can also be changed through /sys/module/overlay/parameters/):

- **”redirect_dir=BOOL”:**
  See OVERLAY_FS_REDIRECT_DIR kernel config option above.
- **”redirect_always_follow=BOOL”:**
  See OVERLAY_FS_REDIRECT_ALWAYS_FOLLOW kernel config option above.
- **”redirect_max=NUM”:**
  The maximum number of bytes in an absolute redirect (default is 256).

Mount options:

- **”redirect_dir=on”:**
  Redirects are enabled.
- **”redirect_dir=follow”:**
  Redirects are not created, but followed.
- **”redirect_dir=nofollow”:**
  Redirects are not created and not followed.

- **”redirect_dir=off”:**
  If ”redirect_always_follow” is enabled in the kernel/module config, this ”off” translates to ”follow”, otherwise it translates to ”nofollow”.

When the NFS export feature is enabled, every copied up directory is indexed by the file handle of the lower inode and a file handle of the upper directory is stored in a ”trusted.overlay.upper” extended attribute on the index entry. On lookup of a merged directory, if the upper directory does not match the file handle stores in the index, that is an indication that multiple upper directories may be redirected to the same lower directory. In that case, lookup returns an error and warns about a possible inconsistency.

Because lower layer redirects cannot be verified with the index, enabling NFS export support on an overlay filesystem with no upper layer requires turning off redirect follow (e.g. ”redirect_dir=nofollow”).

### 3.42.7 Non-directories

Objects that are not directories (files, symlinks, device-special files etc.) are presented either from the upper or lower filesystem as appropriate. When a file in the lower filesystem is accessed in a way the requires write-access, such as opening for write access, changing some metadata etc., the file is first copied from the lower filesystem to the upper filesystem (copy_up). Note that creating a hard-link also requires copy_up, though of course creation of a symlink does not.

The copy_up may turn out to be unnecessary, for example if the file is opened for read-write but the data is not modified.

The copy_up process first makes sure that the containing directory exists in the upper filesystem - creating it and any parents as necessary. It then creates the object with the same metadata (owner, mode, mtime, symlink-target etc.) and then if the object is a file, the data is copied from the lower to the upper filesystem. Finally any extended attributes are copied up.

Once the copy_up is complete, the overlay filesystem simply provides direct access to the newly created file in the upper filesystem - future operations on the file are barely noticed by the overlay filesystem (though an operation on the name of the file such as rename or unlink will of course be noticed and handled).

### 3.42.8 Permission model

Permission checking in the overlay filesystem follows these principles:

1) permission check SHOULD return the same result before and after copy up

2) task creating the overlay mount MUST NOT gain additional privileges

3) non-mounting task MAY gain additional privileges through the overlay, compared to direct access on underlying lower or upper filesystems

This is achieved by performing two permission checks on each access:

a) check if current task is allowed access based on local DAC (owner, group, mode and posix acl), as well as MAC checks

b) check if mounting task would be allowed real operation on lower or upper layer based on underlying filesystem permissions, again including MAC checks

Check (a) ensures consistency (1) since owner, group, mode and posix acls are copied up. On the other hand it can result in server enforced permissions (used by NFS, for example) being ignored (3).

Check (b) ensures that no task gains permissions to underlying layers that the mounting task does not have (2). This also means that it is possible to create setups where the consistency rule (1) does not hold; normally, however, the mounting task will have sufficient privileges to perform all operations.

Another way to demonstrate this model is drawing parallels between:

```
mount -t overlay overlay -olowerdir=/lower,upperdir=/upper,... /merged
```

and:

```
cp -a /lower /upper
mount --bind /upper /merged
```

The resulting access permissions should be the same. The difference is in the time of copy (on-demand vs. up-front).

### 3.42.9 Multiple lower layers

Multiple lower layers can now be given using the colon (":") as a separator character between the directory names. For example:

```
mount -t overlay overlay -olowerdir=/lower1:/lower2:/lower3 /merged
```

As the example shows, "upperdir=" and "workdir=" may be omitted. In that case the overlay will be read-only.

The specified lower directories will be stacked beginning from the rightmost one and going left. In the above example lower1 will be the top, lower2 the middle and lower3 the bottom layer.

Note: directory names containing colons can be provided as lower layer by escaping the colons with a single backslash. For example:

```
mount -t overlay overlay -olowerdir=/a\:lower\:\:dir /merged
```

Since kernel version v6.8, directory names containing colons can also be configured as lower layer using the "lowerdir+" mount options and the fsconfig syscall from new mount api. For example:

```
fsconfig(fs_fd, FSCONFIG_SET_STRING, "lowerdir+", "/a:lower::dir", 0);
```

In the latter case, colons in lower layer directory names will be escaped as an octal characters (072) when displayed in /proc/self/mountinfo.

### 3.42.10 Metadata only copy up

When the ”metacopy” feature is enabled, overlayfs will only copy up metadata (as opposed to whole file), when a metadata specific operation like chown/chmod is performed. Full file will be copied up later when file is opened for WRITE operation.

In other words, this is delayed data copy up operation and data is copied up when there is a need to actually modify data.

There are multiple ways to enable/disable this feature. A config option CONFIG_OVERLAY_FS_METACOPY can be set/unset to enable/disable this feature by default. Or one can enable/disable it at module load time with module parameter metacopy=on/off. Lastly, there is also a per mount option metacopy=on/off to enable/disable this feature per mount.

Do not use metacopy=on with untrusted upper/lower directories. Otherwise it is possible that an attacker can create a handcrafted file with appropriate REDIRECT and METACOPY xattrs, and gain access to file on lower pointed by REDIRECT. This should not be possible on local system as setting ”trusted.” xattrs will require CAP_SYS_ADMIN. But it should be possible for untrusted layers like from a pen drive.

Note: redirect_dir={off|nofollow|follow[*]} and nfs_export=on mount options conflict with metacopy=on, and will result in an error.

[*] redirect_dir=follow only conflicts with metacopy=on if upperdir=... is given.

### 3.42.11 Data-only lower layers

With ”metacopy” feature enabled, an overlayfs regular file may be a composition of information from up to three different layers:

1) metadata from a file in the upper layer

2) st_ino and st_dev object identifier from a file in a lower layer

3) data from a file in another lower layer (further below)

The ”lower data” file can be on any lower layer, except from the top most lower layer.

Below the top most lower layer, any number of lower most layers may be defined as ”data-only” lower layers, using double colon (”::”) separators. A normal lower layer is not allowed to be below a data-only layer, so single colon separators are not allowed to the right of double colon (”::”) separators.

For example:

```
mount -t overlay overlay -olowerdir=/l1:/l2:/l3::/do1::/do2 /merged
```

The paths of files in the ”data-only” lower layers are not visible in the merged overlayfs directories and the metadata and st_ino/st_dev of files in the ”data-only” lower layers are not visible in overlayfs inodes.

Only the data of the files in the ”data-only” lower layers may be visible when a ”metacopy” file in one of the lower layers above it, has a ”redirect” to the absolute path of the ”lower data” file in the ”data-only” lower layer.

Since kernel version v6.8, ”data-only” lower layers can also be added using the ”datadir+” mount options and the fsconfig syscall from new mount api. For example:

```
fsconfig(fs_fd, FSCONFIG_SET_STRING, "lowerdir+", "/l1", 0);
fsconfig(fs_fd, FSCONFIG_SET_STRING, "lowerdir+", "/l2", 0);
fsconfig(fs_fd, FSCONFIG_SET_STRING, "lowerdir+", "/l3", 0);
fsconfig(fs_fd, FSCONFIG_SET_STRING, "datadir+", "/do1", 0);
fsconfig(fs_fd, FSCONFIG_SET_STRING, "datadir+", "/do2", 0);
```

### 3.42.12 fs-verity support

During metadata copy up of a lower file, if the source file has fs-verity enabled and overlay verity support is enabled, then the digest of the lower file is added to the ”trusted.overlay.metacopy” xattr. This is then used to verify the content of the lower file each the time the metacopy file is opened.

When a layer containing verity xattrs is used, it means that any such metacopy file in the upper layer is guaranteed to match the content that was in the lower at the time of the copy-up. If at any time (during a mount, after a remount, etc) such a file in the lower is replaced or modified in any way, access to the corresponding file in overlayfs will result in EIO errors (either on open, due to overlayfs digest check, or from a later read due to fs-verity) and a detailed error is printed to the kernel logs. For more details of how fs-verity file access works, see *Documentation/filesystems/fsverity.rst*.

Verity can be used as a general robustness check to detect accidental changes in the overlayfs directories in use. But, with additional care it can also give more powerful guarantees. For example, if the upper layer is fully trusted (by using dm-verity or something similar), then an untrusted lower layer can be used to supply validated file content for all metacopy files. If additionally the untrusted lower directories are specified as ”Data-only”, then they can only supply such file content, and the entire mount can be trusted to match the upper layer.

This feature is controlled by the ”verity” mount option, which supports these values:

- **”off”:**
  The metacopy digest is never generated or used. This is the default if verity option is not specified.
- **”on”:**
  Whenever a metacopy files specifies an expected digest, the corresponding data file must match the specified digest. When generating a metacopy file the verity digest will be set in it based on the source file (if it has one).
- **”require”:**
  Same as ”on”, but additionally all metacopy files must specify a digest (or EIO is returned on open). This means metadata copy up will only be used if the data file has fs-verity enabled, otherwise a full copy-up is used.

### 3.42.13 Sharing and copying layers

Lower layers may be shared among several overlay mounts and that is indeed a very common practice. An overlay mount may use the same lower layer path as another overlay mount and it may use a lower layer path that is beneath or above the path of another overlay lower layer path.

Using an upper layer path and/or a workdir path that are already used by another overlay mount is not allowed and may fail with EBUSY. Using partially overlapping paths is not allowed and may fail with EBUSY. If files are accessed from two overlayfs mounts which share or overlap the upper layer and/or workdir path the behavior of the overlay is undefined, though it will not result in a crash or deadlock.

Mounting an overlay using an upper layer path, where the upper layer path was previously used by another mounted overlay in combination with a different lower layer path, is allowed, unless the ”index” or ”metacopy” features are enabled.

With the ”index” feature, on the first time mount, an NFS file handle of the lower layer root directory, along with the UUID of the lower filesystem, are encoded and stored in the ”trusted.overlay.origin” extended attribute on the upper layer root directory. On subsequent mount attempts, the lower root directory file handle and lower filesystem UUID are compared to the stored origin in upper root directory. On failure to verify the lower root origin, mount will fail with ESTALE. An overlayfs mount with ”index” enabled will fail with EOPNOTSUPP if the lower filesystem does not support NFS export, lower filesystem does not have a valid UUID or if the upper filesystem does not support extended attributes.

For the ”metacopy” feature, there is no verification mechanism at mount time. So if same upper is mounted with different set of lower, mount probably will succeed but expect the unexpected later on. So don't do it.

It is quite a common practice to copy overlay layers to a different directory tree on the same or different underlying filesystem, and even to a different machine. With the ”index” feature, trying to mount the copied layers will fail the verification of the lower root file handle.

### 3.42.14 Nesting overlayfs mounts

It is possible to use a lower directory that is stored on an overlayfs mount. For regular files this does not need any special care. However, files that have overlayfs attributes, such as whiteouts or ”overlay.*” xattrs will be interpreted by the underlying overlayfs mount and stripped out. In order to allow the second overlayfs mount to see the attributes they must be escaped.

Overlayfs specific xattrs are escaped by using a special prefix of ”overlay.overlay.”. So, a file with a ”trusted.overlay.overlay.metacopy” xattr in the lower dir will be exposed as a regular file with a ”trusted.overlay.metacopy” xattr in the overlayfs mount. This can be nested by repeating the prefix multiple time, as each instance only removes one prefix.

A lower dir with a regular whiteout will always be handled by the overlayfs mount, so to support storing an effective whiteout file in an overlayfs mount an alternative form of whiteout is supported. This form is a regular, zero-size file with the ”overlay.whiteout” xattr set, inside a directory with the ”overlay.opaque” xattr set to ”x” (see *whiteouts and opaque directories*). These alternative whiteouts are never created by overlayfs, but can be used by userspace tools (like containers) that generate lower layers. These alternative whiteouts can be escaped using the standard xattr escape mechanism in order to properly nest to any depth.

### 3.42.15 Non-standard behavior

Current version of overlayfs can act as a mostly POSIX compliant filesystem.

This is the list of cases that overlayfs doesn't currently handle:

a) POSIX mandates updating st_atime for reads. This is currently not done in the case when the file resides on a lower layer.

b) If a file residing on a lower layer is opened for read-only and then memory mapped with MAP_SHARED, then subsequent changes to the file are not reflected in the memory mapping.

c) If a file residing on a lower layer is being executed, then opening that file for write or truncating the file will not be denied with ETXTBSY.

The following options allow overlayfs to act more like a standards compliant filesystem:

#### redirect_dir

Enabled with the mount option or module option: "redirect_dir=on" or with the kernel config option CONFIG_OVERLAY_FS_REDIRECT_DIR=y.

If this feature is disabled, then rename(2) on a lower or merged directory will fail with EXDEV ("Invalid cross-device link").

#### index

Enabled with the mount option or module option "index=on" or with the kernel config option CONFIG_OVERLAY_FS_INDEX=y.

If this feature is disabled and a file with multiple hard links is copied up, then this will "break" the link. Changes will not be propagated to other names referring to the same inode.

#### xino

Enabled with the mount option "xino=auto" or "xino=on", with the module option "xino_auto=on" or with the kernel config option CONFIG_OVERLAY_FS_XINO_AUTO=y. Also implicitly enabled by using the same underlying filesystem for all layers making up the overlay.

If this feature is disabled or the underlying filesystem doesn't have enough free bits in the inode number, then overlayfs will not be able to guarantee that the values of st_ino and st_dev returned by stat(2) and the value of d_ino returned by readdir(3) will act like on a normal filesystem. E.g. the value of st_dev may be different for two objects in the same overlay filesystem and the value of st_ino for filesystem objects may not be persistent and could change even while the overlay filesystem is mounted, as summarized in the *Inode properties* table above.

### 3.42.16 Changes to underlying filesystems

Changes to the underlying filesystems while part of a mounted overlay filesystem are not allowed. If the underlying filesystem is changed, the behavior of the overlay is undefined, though it will not result in a crash or deadlock.

Offline changes, when the overlay is not mounted, are allowed to the upper tree. Offline changes to the lower tree are only allowed if the ”metacopy”, ”index”, ”xino” and ”redirect_dir” features have not been used. If the lower tree is modified and any of these features has been used, the behavior of the overlay is undefined, though it will not result in a crash or deadlock.

When the overlay NFS export feature is enabled, overlay filesystems behavior on offline changes of the underlying lower layer is different than the behavior when NFS export is disabled.

On every copy_up, an NFS file handle of the lower inode, along with the UUID of the lower filesystem, are encoded and stored in an extended attribute ”trusted.overlay.origin” on the upper inode.

When the NFS export feature is enabled, a lookup of a merged directory, that found a lower directory at the lookup path or at the path pointed to by the ”trusted.overlay.redirect” extended attribute, will verify that the found lower directory file handle and lower filesystem UUID match the origin file handle that was stored at copy_up time. If a found lower directory does not match the stored origin, that directory will not be merged with the upper directory.

### 3.42.17 NFS export

When the underlying filesystems supports NFS export and the ”nfs_export” feature is enabled, an overlay filesystem may be exported to NFS.

With the ”nfs_export” feature, on copy_up of any lower object, an index entry is created under the index directory. The index entry name is the hexadecimal representation of the copy up origin file handle. For a non-directory object, the index entry is a hard link to the upper inode. For a directory object, the index entry has an extended attribute ”trusted.overlay.upper” with an encoded file handle of the upper directory inode.

When encoding a file handle from an overlay filesystem object, the following rules apply:

1. For a non-upper object, encode a lower file handle from lower inode
2. For an indexed object, encode a lower file handle from copy_up origin
3. For a pure-upper object and for an existing non-indexed upper object, encode an upper file handle from upper inode

The encoded overlay file handle includes:

- Header including path type information (e.g. lower/upper)
- UUID of the underlying filesystem
- Underlying filesystem encoding of underlying inode

This encoding format is identical to the encoding format file handles that are stored in extended attribute ”trusted.overlay.origin”.

When decoding an overlay file handle, the following steps are followed:

1. Find underlying layer by UUID and path type information.

2. Decode the underlying filesystem file handle to underlying dentry.
3. For a lower file handle, lookup the handle in index directory by name.
4. If a whiteout is found in index, return ESTALE. This represents an overlay object that was deleted after its file handle was encoded.
5. For a non-directory, instantiate a disconnected overlay dentry from the decoded underlying dentry, the path type and index inode, if found.
6. For a directory, use the connected underlying decoded dentry, path type and index, to lookup a connected overlay dentry.

Decoding a non-directory file handle may return a disconnected dentry. copy_up of that disconnected dentry will create an upper index entry with no upper alias.

When overlay filesystem has multiple lower layers, a middle layer directory may have a "redirect" to lower directory. Because middle layer "redirects" are not indexed, a lower file handle that was encoded from the "redirect" origin directory, cannot be used to find the middle or upper layer directory. Similarly, a lower file handle that was encoded from a descendant of the "redirect" origin directory, cannot be used to reconstruct a connected overlay path. To mitigate the cases of directories that cannot be decoded from a lower file handle, these directories are copied up on encode and encoded as an upper file handle. On an overlay filesystem with no upper layer this mitigation cannot be used NFS export in this setup requires turning off redirect follow (e.g. "redirect_dir=nofollow").

The overlay filesystem does not support non-directory connectable file handles, so exporting with the 'subtree_check' exportfs configuration will cause failures to lookup files over NFS.

When the NFS export feature is enabled, all directory index entries are verified on mount time to check that upper file handles are not stale. This verification may cause significant overhead in some cases.

Note: the mount options index=off,nfs_export=on are conflicting for a read-write mount and will result in an error.

Note: the mount option uuid=off can be used to replace UUID of the underlying filesystem in file handles with null, and effectively disable UUID checks. This can be useful in case the underlying disk is copied and the UUID of this copy is changed. This is only applicable if all lower/upper/work directories are on the same filesystem, otherwise it will fallback to normal behaviour.

### 3.42.18 UUID and fsid

The UUID of overlayfs instance itself and the fsid reported by statfs(2) are controlled by the "uuid" mount option, which supports these values:

- **"null":**
  UUID of overlayfs is null. fsid is taken from upper most filesystem.
- **"off":**
  UUID of overlayfs is null. fsid is taken from upper most filesystem. UUID of underlying layers is ignored.
- **"on":**
  UUID of overlayfs is generated and used to report a unique fsid. UUID is stored in

xattr ”trusted.overlay.uuid”, making overlayfs fsid unique and persistent. This option requires an overlayfs with upper filesystem that supports xattrs.

- **”auto”: (default)**
  UUID is taken from xattr ”trusted.overlay.uuid” if it exists. Upgrade to ”uuid=on” on first time mount of new overlay filesystem that meets the prerequites. Downgrade to ”uuid=null” for existing overlay filesystems that were never mounted with ”uuid=on”.

### 3.42.19 Volatile mount

This is enabled with the ”volatile” mount option. Volatile mounts are not guaranteed to survive a crash. It is strongly recommended that volatile mounts are only used if data written to the overlay can be recreated without significant effort.

The advantage of mounting with the ”volatile” option is that all forms of sync calls to the upper filesystem are omitted.

In order to avoid a giving a false sense of safety, the syncfs (and fsync) semantics of volatile mounts are slightly different than that of the rest of VFS. If any writeback error occurs on the upperdir's filesystem after a volatile mount takes place, all sync functions will return an error. Once this condition is reached, the filesystem will not recover, and every subsequent sync call will return an error, even if the upperdir has not experience a new error since the last sync call.

When overlay is mounted with ”volatile” option, the directory ”$workdir/work/incompat/volatile” is created. During next mount, overlay checks for this directory and refuses to mount if present. This is a strong indicator that user should throw away upper and work directories and create fresh one. In very limited cases where the user knows that the system has not crashed and contents of upperdir are intact, The ”volatile” directory can be removed.

### 3.42.20 User xattr

The ”-o userxattr” mount option forces overlayfs to use the ”user.overlay.” xattr namespace instead of ”trusted.overlay.”. This is useful for unprivileged mounting of overlayfs.

### 3.42.21 Testsuite

There's a testsuite originally developed by David Howells and currently maintained by Amir Goldstein at:

https://github.com/amir73il/unionmount-testsuite.git

Run as root:

```
# cd unionmount-testsuite
# ./run --ov --verify
```

# 3.43 The /proc Filesystem

| | | |
|---|---|---|
| /proc/sys | Terrehon Bowden <terrehon@pacbell.net>, Bodo Bauer <bb@ricochet.net> | October 7 1999 |
| 2.4.x update | Jorge Nerin <comandante@zaralinux.com> | November 14 2000 |
| move /proc/sys | Shen Feng <shen@cn.fujitsu.com> | April 1 2009 |
| fixes/update part 1.1 | Stefani Seibold <stefani@seibold.net> | June 9 2009 |

## 3.43.1 Preface

### 0.1 Introduction/Credits

This documentation is part of a soon (or so we hope) to be released book on the SuSE Linux distribution. As there is no complete documentation for the /proc file system and we've used many freely available sources to write these chapters, it seems only fair to give the work back to the Linux community. This work is based on the 2.2.* kernel version and the upcoming 2.4.*. I'm afraid it's still far from complete, but we hope it will be useful. As far as we know, it is the first 'all-in-one' document about the /proc file system. It is focused on the Intel x86 hardware, so if you are looking for PPC, ARM, SPARC, AXP, etc., features, you probably won't find what you are looking for. It also only covers IPv4 networking, not IPv6 nor other protocols - sorry. But additions and patches are welcome and will be added to this document if you mail them to Bodo.

We'd like to thank Alan Cox, Rik van Riel, and Alexey Kuznetsov and a lot of other people for help compiling this documentation. We'd also like to extend a special thank you to Andi Kleen for documentation, which we relied on heavily to create this document, as well as the additional information he provided. Thanks to everybody else who contributed source or docs to the Linux kernel and helped create a great piece of software... :)

If you have any comments, corrections or additions, please don't hesitate to contact Bodo Bauer at bb@ricochet.net. We'll be happy to add them to this document.

The latest version of this document is available online at https://www.kernel.org/doc/html/latest/filesystems/proc.html

If the above direction does not works for you, you could try the kernel mailing list at linux-kernel@vger.kernel.org and/or try to reach me at comandante@zaralinux.com.

### 0.2 Legal Stuff

We don't guarantee the correctness of this document, and if you come to us complaining about how you screwed up your system because of incorrect documentation, we won't feel responsible...

### 3.43.2 Chapter 1: Collecting System Information

#### In This Chapter

- Investigating the properties of the pseudo file system /proc and its ability to provide information on the running Linux system
- Examining /proc's structure
- Uncovering various information about the kernel and the processes running on the system

---

The proc file system acts as an interface to internal data structures in the kernel. It can be used to obtain information about the system and to change certain kernel parameters at runtime (sysctl).

First, we'll take a look at the read-only parts of /proc. In Chapter 2, we show you how you can use /proc/sys to change settings.

#### 1.1 Process-Specific Subdirectories

The directory /proc contains (among other things) one subdirectory for each process running on the system, which is named after the process ID (PID).

The link 'self' points to the process reading the file system. Each process subdirectory has the entries listed in Table 1-1.

Note that an open file descriptor to /proc/<pid> or to any of its contained files or subdirectories does not prevent <pid> being reused for some other process in the event that <pid> exits. Operations on open /proc/<pid> file descriptors corresponding to dead processes never act on any new process that the kernel may, through chance, have also assigned the process ID <pid>. Instead, operations on these FDs usually fail with ESRCH.

Table 6: Table 1-1: Process specific entries in /proc

| File | Content |
|---|---|
| clear_refs | Clears page referenced bits shown in smaps output |
| cmdline | Command line arguments |
| cpu | Current and last cpu in which it was executed (2.4)(smp) |
| cwd | Link to the current working directory |
| environ | Values of environment variables |
| exe | Link to the executable of this process |
| fd | Directory, which contains all file descriptors |
| maps | Memory maps to executables and library files (2.4) |
| mem | Memory held by this process |
| root | Link to the root directory of this process |
| stat | Process status |
| statm | Process memory status information |
| status | Process status in human readable form |
| wchan | Present with CONFIG_KALLSYMS=y: it shows the kernel function symbol the task is blocked in - or "0" if not blocked. |
| pagemap | Page table |
| stack | Report full stack trace, enable via CONFIG_STACKTRACE |
| smaps | An extension based on maps, showing the memory consumption of each mapping and flags associated with it |
| smaps_rollup | Accumulated smaps stats for all mappings of the process. This can be derived from smaps, but is faster and more convenient |
| numa_maps | An extension based on maps, showing the memory locality and binding policy as well as mem usage (in pages) of each mapping. |

For example, to get the status information of a process, all you have to do is read the file /proc/PID/status:

```
>cat /proc/self/status
Name:   cat
State:  R (running)
Tgid:   5452
Pid:    5452
PPid:   743
TracerPid:      0                                                    (2.4)
Uid:    501     501     501     501
Gid:    100     100     100     100
FDSize: 256
Groups: 100 14 16
Kthread:    0
VmPeak:     5004 kB
VmSize:     5004 kB
VmLck:         0 kB
VmHWM:       476 kB
VmRSS:       476 kB
RssAnon:             352 kB
RssFile:             120 kB
RssShmem:              4 kB
VmData:      156 kB
```

```
VmStk:        88 kB
VmExe:        68 kB
VmLib:      1412 kB
VmPTE:        20 kb
VmSwap:        0 kB
HugetlbPages:          0 kB
CoreDumping:    0
THP_enabled:    1
Threads:        1
SigQ:   0/28578
SigPnd: 0000000000000000
ShdPnd: 0000000000000000
SigBlk: 0000000000000000
SigIgn: 0000000000000000
SigCgt: 0000000000000000
CapInh: 00000000fffffeff
CapPrm: 0000000000000000
CapEff: 0000000000000000
CapBnd: ffffffffffffffff
CapAmb: 0000000000000000
NoNewPrivs:     0
Seccomp:        0
Speculation_Store_Bypass:       thread vulnerable
SpeculationIndirectBranch:      conditional enabled
voluntary_ctxt_switches:        0
nonvoluntary_ctxt_switches:     1
```

This shows you nearly the same information you would get if you viewed it with the ps command. In fact, ps uses the proc file system to obtain its information. But you get a more detailed view of the process by reading the file /proc/PID/status. It fields are described in table 1-2.

The statm file contains more detailed information about the process memory usage. Its seven fields are explained in Table 1-3. The stat file contains detailed information about the process itself. Its fields are explained in Table 1-4.

(for SMP CONFIG users)

For making accounting scalable, RSS related information are handled in an asynchronous manner and the value may not be very precise. To see a precise snapshot of a moment, you can see /proc/<pid>/smaps file and scan page table. It's slow but very precise.

Table 7: Table 1-2: Contents of the status fields (as of 4.19)

| Field | Content |
|---|---|
| Name | filename of the executable |
| Umask | file mode creation mask |
| State | state (R is running, S is sleeping, D is sleeping in an uninterruptible w |
| Tgid | thread group ID |
| Ngid | NUMA group ID (0 if none) |
| Pid | process id |

Table 7 – continued from previous page

| Field | Content |
|---|---|
| PPid | process id of the parent process |
| TracerPid | PID of process tracing this process (0 if not, or the tracer is outside of |
| Uid | Real, effective, saved set, and file system UIDs |
| Gid | Real, effective, saved set, and file system GIDs |
| FDSize | number of file descriptor slots currently allocated |
| Groups | supplementary group list |
| NStgid | descendant namespace thread group ID hierarchy |
| NSpid | descendant namespace process ID hierarchy |
| NSpgid | descendant namespace process group ID hierarchy |
| NSsid | descendant namespace session ID hierarchy |
| Kthread | kernel thread flag, 1 is yes, 0 is no |
| VmPeak | peak virtual memory size |
| VmSize | total program size |
| VmLck | locked memory size |
| VmPin | pinned memory size |
| VmHWM | peak resident set size (”high water mark”) |
| VmRSS | size of memory portions. It contains the three following parts (VmRSS |
| RssAnon | size of resident anonymous memory |
| RssFile | size of resident file mappings |
| RssShmem | size of resident shmem memory (includes SysV shm, mapping of tmpfs |
| VmData | size of private data segments |
| VmStk | size of stack segments |
| VmExe | size of text segment |
| VmLib | size of shared library code |
| VmPTE | size of page table entries |
| VmSwap | amount of swap used by anonymous private data (shmem swap usage |
| HugetlbPages | size of hugetlb memory portions |
| CoreDumping | process's memory is currently being dumped (killing the process may |
| THP_enabled | process is allowed to use THP (returns 0 when PR_SET_THP_DISABLE |
| Threads | number of threads |
| SigQ | number of signals queued/max. number for queue |
| SigPnd | bitmap of pending signals for the thread |
| ShdPnd | bitmap of shared pending signals for the process |
| SigBlk | bitmap of blocked signals |
| SigIgn | bitmap of ignored signals |
| SigCgt | bitmap of caught signals |
| CapInh | bitmap of inheritable capabilities |
| CapPrm | bitmap of permitted capabilities |
| CapEff | bitmap of effective capabilities |
| CapBnd | bitmap of capabilities bounding set |
| CapAmb | bitmap of ambient capabilities |
| NoNewPrivs | no_new_privs, like prctl(PR_GET_NO_NEW_PRIV, ...) |
| Seccomp | seccomp mode, like prctl(PR_GET_SECCOMP, ...) |
| Speculation_Store_Bypass | speculative store bypass mitigation status |
| SpeculationIndirectBranch | indirect branch speculation mode |
| Cpus_allowed | mask of CPUs on which this process may run |
| Cpus_allowed_list | Same as previous, but in ”list format” |

Table 7 – continued from previous page

| Field | Content |
|---|---|
| Mems_allowed | mask of memory nodes allowed to this process |
| Mems_allowed_list | Same as previous, but in ”list format” |
| voluntary_ctxt_switches | number of voluntary context switches |
| nonvoluntary_ctxt_switches | number of non voluntary context switches |

Table 8: Table 1-3: Contents of the statm fields (as of 2.6.8-rc3)

| Field | Content | |
|---|---|---|
| size | total program size (pages) | (same as VmSize in status) |
| resident | size of memory portions (pages) | (same as VmRSS in status) |
| shared | number of pages that are shared | (i.e. backed by a file, same as RssFile+RssShmem in status) |
| trs | number of pages that are 'code' | (not including libs; broken, includes data segment) |
| lrs | number of pages of library | (always 0 on 2.6) |
| drs | number of pages of data/stack | (including libs; broken, includes library text) |
| dt | number of dirty pages | (always 0 on 2.6) |

Table 9: Table 1-4: Contents of the stat fields (as of 2.6.30-rc7)

| Field | Content |
|---|---|
| pid | process id |
| tcomm | filename of the executable |
| state | state (R is running, S is sleeping, D is sleeping in an uninterruptible wait, Z is zombi |
| ppid | process id of the parent process |
| pgrp | pgrp of the process |
| sid | session id |
| tty_nr | tty the process uses |
| tty_pgrp | pgrp of the tty |
| flags | task flags |
| min_flt | number of minor faults |
| cmin_flt | number of minor faults with child's |
| maj_flt | number of major faults |
| cmaj_flt | number of major faults with child's |
| utime | user mode jiffies |
| stime | kernel mode jiffies |
| cutime | user mode jiffies with child's |
| cstime | kernel mode jiffies with child's |
| priority | priority level |
| nice | nice level |
| num_threads | number of threads |
| it_real_value | (obsolete, always 0) |
| start_time | time the process started after system boot |

Table 9 – continued from previous page

| Field | Content |
|---|---|
| vsize | virtual memory size |
| rss | resident set memory size |
| rsslim | current limit in bytes on the rss |
| start_code | address above which program text can run |
| end_code | address below which program text can run |
| start_stack | address of the start of the main process stack |
| esp | current value of ESP |
| eip | current value of EIP |
| pending | bitmap of pending signals |
| blocked | bitmap of blocked signals |
| sigign | bitmap of ignored signals |
| sigcatch | bitmap of caught signals |
| 0 | (place holder, used to be the wchan address, use /proc/PID/wchan instead) |
| 0 | (place holder) |
| 0 | (place holder) |
| exit_signal | signal to send to parent thread on exit |
| task_cpu | which CPU the task is scheduled on |
| rt_priority | realtime priority |
| policy | scheduling policy (man sched_setscheduler) |
| blkio_ticks | time spent waiting for block IO |
| gtime | guest time of the task in jiffies |
| cgtime | guest time of the task children in jiffies |
| start_data | address above which program data+bss is placed |
| end_data | address below which program data+bss is placed |
| start_brk | address above which program heap can be expanded with brk() |
| arg_start | address above which program command line is placed |
| arg_end | address below which program command line is placed |
| env_start | address above which program environment is placed |
| env_end | address below which program environment is placed |
| exit_code | the thread's exit_code in the form reported by the waitpid system call |

The /proc/PID/maps file contains the currently mapped memory regions and their access permissions.

The format is:

```
address           perms offset  dev   inode      pathname

08048000-08049000 r-xp 00000000 03:00 8312       /opt/test
08049000-0804a000 rw-p 00001000 03:00 8312       /opt/test
0804a000-0806b000 rw-p 00000000 00:00 0          [heap]
a7cb1000-a7cb2000 ---p 00000000 00:00 0
a7cb2000-a7eb2000 rw-p 00000000 00:00 0
a7eb2000-a7eb3000 ---p 00000000 00:00 0
a7eb3000-a7ed5000 rw-p 00000000 00:00 0
a7ed5000-a8008000 r-xp 00000000 03:00 4222       /lib/libc.so.6
a8008000-a800a000 r--p 00133000 03:00 4222       /lib/libc.so.6
a800a000-a800b000 rw-p 00135000 03:00 4222       /lib/libc.so.6
a800b000-a800e000 rw-p 00000000 00:00 0
```

```
a800e000-a8022000 r-xp 00000000 03:00 14462      /lib/libpthread.so.0
a8022000-a8023000 r--p 00013000 03:00 14462      /lib/libpthread.so.0
a8023000-a8024000 rw-p 00014000 03:00 14462      /lib/libpthread.so.0
a8024000-a8027000 rw-p 00000000 00:00 0
a8027000-a8043000 r-xp 00000000 03:00 8317       /lib/ld-linux.so.2
a8043000-a8044000 r--p 0001b000 03:00 8317       /lib/ld-linux.so.2
a8044000-a8045000 rw-p 0001c000 03:00 8317       /lib/ld-linux.so.2
aff35000-aff4a000 rw-p 00000000 00:00 0          [stack]
ffffe000-fffff000 r-xp 00000000 00:00 0          [vdso]
```

where ”address” is the address space in the process that it occupies, ”perms” is a set of permissions:

```
r = read
w = write
x = execute
s = shared
p = private (copy on write)
```

”offset” is the offset into the mapping, ”dev” is the device (major:minor), and ”inode” is the inode on that device. 0 indicates that no inode is associated with the memory region, as the case would be with BSS (uninitialized data). The ”pathname” shows the name associated file for this mapping. If the mapping is not associated with a file:

| | |
|---|---|
| [heap] | the heap of the program |
| [stack] | the stack of the main process |
| [vdso] | the ”virtual dynamic shared object”, the kernel system call handler |
| [anon:<name>] | a private anonymous mapping that has been named by userspace |
| [anon_shmem:<name>] | an anonymous shared memory mapping that has been named by userspace |

or if empty, the mapping is anonymous.

The /proc/PID/smaps is an extension based on maps, showing the memory consumption for each of the process's mappings. For each mapping (aka Virtual Memory Area, or VMA) there is a series of lines such as the following:

```
08048000-080bc000 r-xp 00000000 03:02 13130      /bin/bash

Size:               1084 kB
KernelPageSize:        4 kB
MMUPageSize:           4 kB
Rss:                 892 kB
Pss:                 374 kB
Pss_Dirty:             0 kB
Shared_Clean:        892 kB
Shared_Dirty:          0 kB
Private_Clean:         0 kB
```

```
Private_Dirty:         0 kB
Referenced:          892 kB
Anonymous:             0 kB
KSM:                   0 kB
LazyFree:              0 kB
AnonHugePages:         0 kB
ShmemPmdMapped:        0 kB
Shared_Hugetlb:        0 kB
Private_Hugetlb:       0 kB
Swap:                  0 kB
SwapPss:               0 kB
KernelPageSize:        4 kB
MMUPageSize:           4 kB
Locked:                0 kB
THPeligible:           0
VmFlags: rd ex mr mw me dw
```

The first of these lines shows the same information as is displayed for the mapping in /proc/PID/maps. Following lines show the size of the mapping (size); the size of each page allocated when backing a VMA (KernelPageSize), which is usually the same as the size in the page table entries; the page size used by the MMU when backing a VMA (in most cases, the same as KernelPageSize); the amount of the mapping that is currently resident in RAM (RSS); the process' proportional share of this mapping (PSS); and the number of clean and dirty shared and private pages in the mapping.

The ”proportional set size” (PSS) of a process is the count of pages it has in memory, where each page is divided by the number of processes sharing it. So if a process has 1000 pages all to itself, and 1000 shared with one other process, its PSS will be 1500. ”Pss_Dirty” is the portion of PSS which consists of dirty pages. (”Pss_Clean” is not included, but it can be calculated by subtracting ”Pss_Dirty” from ”Pss”.)

Note that even a page which is part of a MAP_SHARED mapping, but has only a single pte mapped, i.e. is currently used by only one process, is accounted as private and not as shared.

”Referenced” indicates the amount of memory currently marked as referenced or accessed.

”Anonymous” shows the amount of memory that does not belong to any file. Even a mapping associated with a file may contain anonymous pages: when MAP_PRIVATE and a page is modified, the file page is replaced by a private anonymous copy.

”KSM” reports how many of the pages are KSM pages. Note that KSM-placed zeropages are not included, only actual KSM pages.

”LazyFree” shows the amount of memory which is marked by madvise(MADV_FREE). The memory isn't freed immediately with madvise(). It's freed in memory pressure if the memory is clean. Please note that the printed value might be lower than the real value due to optimizations used in the current implementation. If this is not desirable please file a bug report.

”AnonHugePages” shows the amount of memory backed by transparent hugepage.

”ShmemPmdMapped” shows the amount of shared (shmem/tmpfs) memory backed by huge pages.

”Shared_Hugetlb” and ”Private_Hugetlb” show the amounts of memory backed by hugetlbfs page which is *not* counted in ”RSS” or ”PSS” field for historical reasons. And these are not

included in {Shared,Private}_{Clean,Dirty} field.

”Swap” shows how much would-be-anonymous memory is also used, but out on swap.

For shmem mappings, ”Swap” includes also the size of the mapped (and not replaced by copy-on-write) part of the underlying shmem object out on swap. ”SwapPss” shows proportional swap share of this mapping. Unlike ”Swap”, this does not take into account swapped out page of underlying shmem objects. ”Locked” indicates whether the mapping is locked in memory or not.

”THPeligible” indicates whether the mapping is eligible for allocating naturally aligned THP pages of any currently enabled size. 1 if true, 0 otherwise.

”VmFlags” field deserves a separate description. This member represents the kernel flags associated with the particular virtual memory area in two letter encoded manner. The codes are the following:

| | |
|---|---|
| rd | readable |
| wr | writeable |
| ex | executable |
| sh | shared |
| mr | may read |
| mw | may write |
| me | may execute |
| ms | may share |
| gd | stack segment growns down |
| pf | pure PFN range |
| dw | disabled write to the mapped file |
| lo | pages are locked in memory |
| io | memory mapped I/O area |
| sr | sequential read advise provided |
| rr | random read advise provided |
| dc | do not copy area on fork |
| de | do not expand area on remapping |
| ac | area is accountable |
| nr | swap space is not reserved for the area |
| ht | area uses huge tlb pages |
| sf | synchronous page fault |
| ar | architecture specific flag |
| wf | wipe on fork |
| dd | do not include area into core dump |
| sd | soft dirty flag |
| mm | mixed map area |
| hg | huge page advise flag |
| nh | no huge page advise flag |
| mg | mergeable advise flag |
| bt | arm64 BTI guarded page |
| mt | arm64 MTE allocation tags are enabled |
| um | userfaultfd missing tracking |
| uw | userfaultfd wr-protect tracking |
| ss | shadow stack page |

Note that there is no guarantee that every flag and associated mnemonic will be present in all further kernel releases. Things get changed, the flags may be vanished or the reverse -- new added. Interpretation of their meaning might change in future as well. So each consumer of these flags has to follow each specific kernel version for the exact semantic.

This file is only present if the CONFIG_MMU kernel configuration option is enabled.

Note: reading /proc/PID/maps or /proc/PID/smaps is inherently racy (consistent output can be achieved only in the single read call).

This typically manifests when doing partial reads of these files while the memory map is being modified. Despite the races, we do provide the following guarantees:

1) The mapped addresses never go backwards, which implies no two regions will ever overlap.
2) If there is something at a given vaddr during the entirety of the life of the smaps/maps walk, there will be some output for it.

The /proc/PID/smaps_rollup file includes the same fields as /proc/PID/smaps, but their values are the sums of the corresponding values for all mappings of the process. Additionally, it contains these fields:

- Pss_Anon
- Pss_File
- Pss_Shmem

They represent the proportional shares of anonymous, file, and shmem pages, as described for smaps above. These fields are omitted in smaps since each mapping identifies the type (anon, file, or shmem) of all pages it contains. Thus all information in smaps_rollup can be derived from smaps, but at a significantly higher cost.

The /proc/PID/clear_refs is used to reset the PG_Referenced and ACCESSED/YOUNG bits on both physical and virtual pages associated with a process, and the soft-dirty bit on pte (see Documentation/admin-guide/mm/soft-dirty.rst for details). To clear the bits for all the pages associated with the process:

```
> echo 1 > /proc/PID/clear_refs
```

To clear the bits for the anonymous pages associated with the process:

```
> echo 2 > /proc/PID/clear_refs
```

To clear the bits for the file mapped pages associated with the process:

```
> echo 3 > /proc/PID/clear_refs
```

To clear the soft-dirty bit:

```
> echo 4 > /proc/PID/clear_refs
```

To reset the peak resident set size (”high water mark”) to the process's current value:

```
> echo 5 > /proc/PID/clear_refs
```

Any other value written to /proc/PID/clear_refs will have no effect.

The /proc/pid/pagemap gives the PFN, which can be used to find the pageflags using /proc/kpageflags and number of times a page is mapped using /proc/kpagecount. For detailed explanation, see Documentation/admin-guide/mm/pagemap.rst.

The /proc/pid/numa_maps is an extension based on maps, showing the memory locality and binding policy, as well as the memory usage (in pages) of each mapping. The output follows a general format where mapping details get summarized separated by blank spaces, one mapping per each file line:

```
address   policy    mapping details

00400000 default file=/usr/local/bin/app mapped=1 active=0 N3=1 kernelpagesize_
↪kB=4
00600000 default file=/usr/local/bin/app anon=1 dirty=1 N3=1 kernelpagesize_
↪kB=4
3206000000 default file=/lib64/ld-2.12.so mapped=26 mapmax=6 N0=24 N3=2␣
↪kernelpagesize_kB=4
320621f000 default file=/lib64/ld-2.12.so anon=1 dirty=1 N3=1 kernelpagesize_
↪kB=4
3206220000 default file=/lib64/ld-2.12.so anon=1 dirty=1 N3=1 kernelpagesize_
↪kB=4
3206221000 default anon=1 dirty=1 N3=1 kernelpagesize_kB=4
3206800000 default file=/lib64/libc-2.12.so mapped=59 mapmax=21 active=55␣
↪N0=41 N3=18 kernelpagesize_kB=4
320698b000 default file=/lib64/libc-2.12.so
3206b8a000 default file=/lib64/libc-2.12.so anon=2 dirty=2 N3=2 kernelpagesize_
↪kB=4
3206b8e000 default file=/lib64/libc-2.12.so anon=1 dirty=1 N3=1 kernelpagesize_
↪kB=4
3206b8f000 default anon=3 dirty=3 active=1 N3=3 kernelpagesize_kB=4
7f4dc10a2000 default anon=3 dirty=3 N3=3 kernelpagesize_kB=4
7f4dc10b4000 default anon=2 dirty=2 active=1 N3=2 kernelpagesize_kB=4
7f4dc1200000 default file=/anon_hugepage\040(deleted) huge anon=1 dirty=1 N3=1␣
↪kernelpagesize_kB=2048
7fff335f0000 default stack anon=3 dirty=3 N3=3 kernelpagesize_kB=4
7fff3369d000 default mapped=1 mapmax=35 active=0 N3=1 kernelpagesize_kB=4
```

Where:

”address” is the starting address for the mapping;

”policy” reports the NUMA memory policy set for the mapping (see Documentation/admin-guide/mm/numa_memory_policy.rst);

”mapping details” summarizes mapping data such as mapping type, page usage counters, node locality page counters (N0 == node0, N1 == node1, ...) and the kernel page size, in KB, that is backing the mapping up.

### 1.2 Kernel data

Similar to the process entries, the kernel data files give information about the running kernel. The files used to obtain this information are contained in /proc and are listed in Table 1-5. Not all of these will be present in your system. It depends on the kernel configuration and the loaded modules, which files are there, and which are missing.

Table 11: Table 1-5: Kernel info in /proc

| File | Content |
|---|---|
| apm | Advanced power management info |
| bootconfig | Kernel command line obtained from boot config, and, if there were kernel parameters from the boot loader, a "# Parameters from bootloader:" line followed by a line containing those parameters prefixed by "# ". (5.5) |
| buddyinfo | Kernel memory allocator information (see text) (2.5) |
| bus | Directory containing bus specific information |
| cmdline | Kernel command line, both from bootloader and embedded in the kernel image |
| cpuinfo | Info about the CPU |
| devices | Available devices (block and character) |
| dma | Used DMS channels |
| filesystems | Supported filesystems |
| driver | Various drivers grouped here, currently rtc (2.4) |
| execdomains | Execdomains, related to security (2.4) |
| fb | Frame Buffer devices (2.4) |
| fs | File system parameters, currently nfs/exports (2.4) |
| ide | Directory containing info about the IDE subsystem |
| interrupts | Interrupt usage |
| iomem | Memory map (2.4) |
| ioports | I/O port usage |
| irq | Masks for irq to cpu affinity (2.4)(smp?) |
| isapnp | ISA PnP (Plug&Play) Info (2.4) |
| kcore | Kernel core image (can be ELF or A.OUT(deprecated in 2.4)) |
| kmsg | Kernel messages |
| ksyms | Kernel symbol table |

continues on next page

Table 11 – continued from previous page

| File | Content |
|---|---|
| loadavg | **Load average of last 1, 5 & 15 minutes;** number of processes currently runnable (running or on ready queue); total number of processes in system; last pid created. All fields are separated by one space except ”number of processes currently runnable” and ”total number of processes in system”, which are separated by a slash ('/'). Example: 0.61 0.61 0.55 3/828 22084 |
| locks | Kernel locks |
| meminfo | Memory info |
| misc | Miscellaneous |
| modules | List of loaded modules |
| mounts | Mounted filesystems |
| net | Networking info (see text) |
| pagetypeinfo | Additional page allocator information (see text) (2.5) |
| partitions | Table of partitions known to the system |
| pci | Deprecated info of PCI bus (new way -> /proc/bus/pci/, decoupled by lspci (2.4) |
| rtc | Real time clock |
| scsi | SCSI info (see text) |
| slabinfo | Slab pool info |
| softirqs | softirq usage |
| stat | Overall statistics |
| swaps | Swap space utilization |
| sys | See chapter 2 |
| sysvipc | Info of SysVIPC Resources (msg, sem, shm) (2.4) |
| tty | Info of tty drivers |
| uptime | Wall clock since boot, combined idle time of all cpus |
| version | Kernel version |
| video | bttv info of video resources (2.4) |
| vmallocinfo | Show vmalloced areas |

You can, for example, check which interrupts are currently in use and what they are used for by looking in the file /proc/interrupts:

```
> cat /proc/interrupts
           CPU0
  0:    8728810          XT-PIC  timer
  1:        895          XT-PIC  keyboard
  2:          0          XT-PIC  cascade
  3:     531695          XT-PIC  aha152x
  4:    2014133          XT-PIC  serial
```

```
  5:      44401          XT-PIC  pcnet_cs
  8:          2          XT-PIC  rtc
 11:          8          XT-PIC  i82365
 12:     182918          XT-PIC  PS/2 Mouse
 13:          1          XT-PIC  fpu
 14:    1232265          XT-PIC  ide0
 15:          7          XT-PIC  ide1
NMI:          0
```

In 2.4.* a couple of lines where added to this file LOC & ERR (this time is the output of a SMP machine):

```
> cat /proc/interrupts

           CPU0       CPU1
  0:    1243498    1214548    IO-APIC-edge  timer
  1:       8949       8958    IO-APIC-edge  keyboard
  2:          0          0          XT-PIC  cascade
  5:      11286      10161    IO-APIC-edge  soundblaster
  8:          1          0    IO-APIC-edge  rtc
  9:      27422      27407    IO-APIC-edge  3c503
 12:     113645     113873    IO-APIC-edge  PS/2 Mouse
 13:          0          0          XT-PIC  fpu
 14:      22491      24012    IO-APIC-edge  ide0
 15:       2183       2415    IO-APIC-edge  ide1
 17:      30564      30414   IO-APIC-level  eth0
 18:        177        164   IO-APIC-level  bttv
NMI:    2457961    2457959
LOC:    2457882    2457881
ERR:       2155
```

NMI is incremented in this case because every timer interrupt generates a NMI (Non Maskable Interrupt) which is used by the NMI Watchdog to detect lockups.

LOC is the local interrupt counter of the internal APIC of every CPU.

ERR is incremented in the case of errors in the IO-APIC bus (the bus that connects the CPUs in a SMP system. This means that an error has been detected, the IO-APIC automatically retry the transmission, so it should not be a big problem, but you should read the SMP-FAQ.

In 2.6.2* /proc/interrupts was expanded again. This time the goal was for /proc/interrupts to display every IRQ vector in use by the system, not just those considered 'most important'. The new vectors are:

**THR**
: interrupt raised when a machine check threshold counter (typically counting ECC corrected errors of memory or cache) exceeds a configurable threshold. Only available on some systems.

**TRM**
: a thermal event interrupt occurs when a temperature threshold has been exceeded for the CPU. This interrupt may also be generated when the temperature drops back to normal.

**SPU**

a spurious interrupt is some interrupt that was raised then lowered by some IO device before it could be fully processed by the APIC. Hence the APIC sees the interrupt but does not know what device it came from. For this case the APIC will generate the interrupt with a IRQ vector of 0xff. This might also be generated by chipset bugs.

**RES, CAL, TLB**
rescheduling, call and TLB flush interrupts are sent from one CPU to another per the needs of the OS. Typically, their statistics are used by kernel developers and interested users to determine the occurrence of interrupts of the given type.

The above IRQ vectors are displayed only when relevant. For example, the threshold vector does not exist on x86_64 platforms. Others are suppressed when the system is a uniprocessor. As of this writing, only i386 and x86_64 platforms support the new IRQ vector displays.

Of some interest is the introduction of the /proc/irq directory to 2.4. It could be used to set IRQ to CPU affinity. This means that you can ”hook” an IRQ to only one CPU, or to exclude a CPU of handling IRQs. The contents of the irq subdir is one subdir for each IRQ, and two files; default_smp_affinity and prof_cpu_mask.

For example:

```
> ls /proc/irq/
0  10  12  14  16  18  2  4  6  8  prof_cpu_mask
1  11  13  15  17  19  3  5  7  9  default_smp_affinity
> ls /proc/irq/0/
smp_affinity
```

smp_affinity is a bitmask, in which you can specify which CPUs can handle the IRQ. You can set it by doing:

```
> echo 1 > /proc/irq/10/smp_affinity
```

This means that only the first CPU will handle the IRQ, but you can also echo 5 which means that only the first and third CPU can handle the IRQ.

The contents of each smp_affinity file is the same by default:

```
> cat /proc/irq/0/smp_affinity
ffffffff
```

There is an alternate interface, smp_affinity_list which allows specifying a CPU range instead of a bitmask:

```
> cat /proc/irq/0/smp_affinity_list
1024-1031
```

The default_smp_affinity mask applies to all non-active IRQs, which are the IRQs which have not yet been allocated/activated, and hence which lack a /proc/irq/[0-9]* directory.

The node file on an SMP system shows the node to which the device using the IRQ reports itself as being attached. This hardware locality information does not include information about any possible driver locality preference.

prof_cpu_mask specifies which CPUs are to be profiled by the system wide profiler. Default value is ffffffff (all CPUs if there are only 32 of them).

The way IRQs are routed is handled by the IO-APIC, and it's Round Robin between all the CPUs which are allowed to handle it. As usual the kernel has more info than you and does a better job than you, so the defaults are the best choice for almost everyone. [Note this applies only to those IO-APIC's that support "Round Robin" interrupt distribution.]

There are three more important subdirectories in /proc: net, scsi, and sys. The general rule is that the contents, or even the existence of these directories, depend on your kernel configuration. If SCSI is not enabled, the directory scsi may not exist. The same is true with the net, which is there only when networking support is present in the running kernel.

The slabinfo file gives information about memory usage at the slab level. Linux uses slab pools for memory management above page level in version 2.2. Commonly used objects have their own slab pool (such as network buffers, directory cache, and so on).

```
> cat /proc/buddyinfo

Node 0, zone      DMA      0      4      5      4      4      3 ...
Node 0, zone   Normal      1      0      0      1    101      8 ...
Node 0, zone  HighMem      2      0      0      1      1      0 ...
```

External fragmentation is a problem under some workloads, and buddyinfo is a useful tool for helping diagnose these problems. Buddyinfo will give you a clue as to how big an area you can safely allocate, or why a previous allocation failed.

Each column represents the number of pages of a certain order which are available. In this case, there are 0 chunks of 2^0*PAGE_SIZE available in ZONE_DMA, 4 chunks of 2^1*PAGE_SIZE in ZONE_DMA, 101 chunks of 2^4*PAGE_SIZE available in ZONE_NORMAL, etc...

More information relevant to external fragmentation can be found in pagetypeinfo:

```
> cat /proc/pagetypeinfo
Page block order: 9
Pages per block:  512

Free pages count per migrate type at order       0      1      2      3      4␣
 ↪      5      6      7      8      9     10
Node    0, zone      DMA, type    Unmovable      0      0      0      1      1␣
 ↪      1      1      1      1      1      0
Node    0, zone      DMA, type  Reclaimable      0      0      0      0      0␣
 ↪      0      0      0      0      0      0
Node    0, zone      DMA, type      Movable      1      1      2      1      2␣
 ↪      1      1      0      1      0      2
Node    0, zone      DMA, type      Reserve      0      0      0      0      0␣
 ↪      0      0      0      0      1      0
Node    0, zone      DMA, type      Isolate      0      0      0      0      0␣
 ↪      0      0      0      0      0      0
Node    0, zone    DMA32, type    Unmovable    103     54     77      1      1␣
 ↪      1     11      8      7      1      9
Node    0, zone    DMA32, type  Reclaimable      0      0      2      1      0␣
 ↪      0      0      0      1      0      0
Node    0, zone    DMA32, type      Movable    169    152    113     91     77␣
 ↪     54     39     13      6      1    452
Node    0, zone    DMA32, type      Reserve      1      2      2      2      2␣
 ↪      0      1      1      1      1      0
```

```
Node    0, zone    DMA32, type      Isolate      0      0      0      0      0␣
 ↪     0      0      0      0      0      0

Number of blocks type     Unmovable  Reclaimable      Movable      Reserve      ␣
 ↪ Isolate
Node 0, zone      DMA            2            0            5            1      ␣
 ↪      0
Node 0, zone    DMA32           41            6          967            2      ␣
 ↪      0
```

Fragmentation avoidance in the kernel works by grouping pages of different migrate types into the same contiguous regions of memory called page blocks. A page block is typically the size of the default hugepage size, e.g. 2MB on X86-64. By keeping pages grouped based on their ability to move, the kernel can reclaim pages within a page block to satisfy a high-order allocation.

The pagetypinfo begins with information on the size of a page block. It then gives the same type of information as buddyinfo except broken down by migrate-type and finishes with details on how many page blocks of each type exist.

If min_free_kbytes has been tuned correctly (recommendations made by hugeadm from libhugetlbfs https://github.com/libhugetlbfs/libhugetlbfs/), one can make an estimate of the likely number of huge pages that can be allocated at a given point in time. All the "Movable" blocks should be allocatable unless memory has been mlock()'d. Some of the Reclaimable blocks should also be allocatable although a lot of filesystem metadata may have to be reclaimed to achieve this.

### meminfo

Provides information about distribution and utilization of memory. This varies by architecture and compile options. Some of the counters reported here overlap. The memory reported by the non overlapping counters may not add up to the overall memory usage and the difference for some workloads can be substantial. In many cases there are other means to find out additional memory using subsystem specific interfaces, for instance /proc/net/sockstat for TCP memory allocations.

Example output. You may not have all of these fields.

```
> cat /proc/meminfo

MemTotal:       32858820 kB
MemFree:        21001236 kB
MemAvailable:   27214312 kB
Buffers:          581092 kB
Cached:          5587612 kB
SwapCached:            0 kB
Active:          3237152 kB
Inactive:        7586256 kB
Active(anon):      94064 kB
Inactive(anon):  4570616 kB
Active(file):    3143088 kB
Inactive(file):  3015640 kB
```

```
Unevictable:           0 kB
Mlocked:               0 kB
SwapTotal:             0 kB
SwapFree:              0 kB
Zswap:              1904 kB
Zswapped:           7792 kB
Dirty:                12 kB
Writeback:             0 kB
AnonPages:       4654780 kB
Mapped:           266244 kB
Shmem:              9976 kB
KReclaimable:     517708 kB
Slab:             660044 kB
SReclaimable:     517708 kB
SUnreclaim:       142336 kB
KernelStack:       11168 kB
PageTables:        20540 kB
SecPageTables:         0 kB
NFS_Unstable:          0 kB
Bounce:                0 kB
WritebackTmp:          0 kB
CommitLimit:    16429408 kB
Committed_AS:    7715148 kB
VmallocTotal:   34359738367 kB
VmallocUsed:       40444 kB
VmallocChunk:          0 kB
Percpu:            29312 kB
EarlyMemtestBad:       0 kB
HardwareCorrupted:     0 kB
AnonHugePages:   4149248 kB
ShmemHugePages:        0 kB
ShmemPmdMapped:        0 kB
FileHugePages:         0 kB
FilePmdMapped:         0 kB
CmaTotal:              0 kB
CmaFree:               0 kB
HugePages_Total:       0
HugePages_Free:        0
HugePages_Rsvd:        0
HugePages_Surp:        0
Hugepagesize:       2048 kB
Hugetlb:               0 kB
DirectMap4k:      401152 kB
DirectMap2M:    10008576 kB
DirectMap1G:    24117248 kB
```

**MemTotal**
Total usable RAM (i.e. physical RAM minus a few reserved bits and the kernel binary code)

**MemFree**
Total free RAM. On highmem systems, the sum of LowFree+HighFree

**MemAvailable**
An estimate of how much memory is available for starting new applications, without swapping. Calculated from MemFree, SReclaimable, the size of the file LRU lists, and the low watermarks in each zone. The estimate takes into account that the system needs some page cache to function well, and that not all reclaimable slab will be reclaimable, due to items being in use. The impact of those factors will vary from system to system.

**Buffers**
Relatively temporary storage for raw disk blocks shouldn't get tremendously large (20MB or so)

**Cached**
In-memory cache for files read from the disk (the pagecache) as well as tmpfs & shmem. Doesn't include SwapCached.

**SwapCached**
Memory that once was swapped out, is swapped back in but still also is in the swapfile (if memory is needed it doesn't need to be swapped out AGAIN because it is already in the swapfile. This saves I/O)

**Active**
Memory that has been used more recently and usually not reclaimed unless absolutely necessary.

**Inactive**
Memory which has been less recently used. It is more eligible to be reclaimed for other purposes

**Unevictable**
Memory allocated for userspace which cannot be reclaimed, such as mlocked pages, ramfs backing pages, secret memfd pages etc.

**Mlocked**
Memory locked with mlock().

**HighTotal, HighFree**
Highmem is all memory above ~860MB of physical memory. Highmem areas are for use by userspace programs, or for the pagecache. The kernel must use tricks to access this memory, making it slower to access than lowmem.

**LowTotal, LowFree**
Lowmem is memory which can be used for everything that highmem can be used for, but it is also available for the kernel's use for its own data structures. Among many other things, it is where everything from the Slab is allocated. Bad things happen when you're out of lowmem.

**SwapTotal**
total amount of swap space available

**SwapFree**
Memory which has been evicted from RAM, and is temporarily on the disk

**Zswap**
Memory consumed by the zswap backend (compressed size)

**Zswapped**
Amount of anonymous memory stored in zswap (original size)

**Dirty**
Memory which is waiting to get written back to the disk

**Writeback**
Memory which is actively being written back to the disk

**AnonPages**
Non-file backed pages mapped into userspace page tables

**Mapped**
files which have been mmapped, such as libraries

**Shmem**
Total memory used by shared memory (shmem) and tmpfs

**KReclaimable**
Kernel allocations that the kernel will attempt to reclaim under memory pressure. Includes SReclaimable (below), and other direct allocations with a shrinker.

**Slab**
in-kernel data structures cache

**SReclaimable**
Part of Slab, that might be reclaimed, such as caches

**SUnreclaim**
Part of Slab, that cannot be reclaimed on memory pressure

**KernelStack**
Memory consumed by the kernel stacks of all tasks

**PageTables**
Memory consumed by userspace page tables

**SecPageTables**
Memory consumed by secondary page tables, this currently currently includes KVM mmu allocations on x86 and arm64.

**NFS_Unstable**
Always zero. Previous counted pages which had been written to the server, but has not been committed to stable storage.

**Bounce**
Memory used for block device ”bounce buffers”

**WritebackTmp**
Memory used by FUSE for temporary writeback buffers

**CommitLimit**
Based on the overcommit ratio ('vm.overcommit_ratio'), this is the total amount of memory currently available to be allocated on the system. This limit is only adhered to if strict overcommit accounting is enabled (mode 2 in 'vm.overcommit_memory').

The CommitLimit is calculated with the following formula:

```
CommitLimit = ([total RAM pages] - [total huge TLB pages]) *
               overcommit_ratio / 100 + [total swap pages]
```

For example, on a system with 1G of physical RAM and 7G of swap with a *vm.overcommit_ratio* of 30 it would yield a CommitLimit of 7.3G.

For more details, see the memory overcommit documentation in mm/overcommit-accounting.

**Committed_AS**
The amount of memory presently allocated on the system. The committed memory is a sum of all of the memory which has been allocated by processes, even if it has not been "used" by them as of yet. A process which malloc()'s 1G of memory, but only touches 300M of it will show up as using 1G. This 1G is memory which has been "committed" to by the VM and can be used at any time by the allocating application. With strict overcommit enabled on the system (mode 2 in 'vm.overcommit_memory'), allocations which would exceed the CommitLimit (detailed above) will not be permitted. This is useful if one needs to guarantee that processes will not fail due to lack of memory once that memory has been successfully allocated.

**VmallocTotal**
total size of vmalloc virtual address space

**VmallocUsed**
amount of vmalloc area which is used

**VmallocChunk**
largest contiguous block of vmalloc area which is free

**Percpu**
Memory allocated to the percpu allocator used to back percpu allocations. This stat excludes the cost of metadata.

**EarlyMemtestBad**
The amount of RAM/memory in kB, that was identified as corrupted by early memtest. If memtest was not run, this field will not be displayed at all. Size is never rounded down to 0 kB. That means if 0 kB is reported, you can safely assume there was at least one pass of memtest and none of the passes found a single faulty byte of RAM.

**HardwareCorrupted**
The amount of RAM/memory in KB, the kernel identifies as corrupted.

**AnonHugePages**
Non-file backed huge pages mapped into userspace page tables

**ShmemHugePages**
Memory used by shared memory (shmem) and tmpfs allocated with huge pages

**ShmemPmdMapped**
Shared memory mapped into userspace with huge pages

**FileHugePages**
Memory used for filesystem data (page cache) allocated with huge pages

**FilePmdMapped**
Page cache mapped into userspace with huge pages

**CmaTotal**
Memory reserved for the Contiguous Memory Allocator (CMA)

**CmaFree**
Free remaining memory in the CMA reserves

**HugePages_Total, HugePages_Free, HugePages_Rsvd, HugePages_Surp, Hugepagesize, Hugetlb**
See Documentation/admin-guide/mm/hugetlbpage.rst.

**DirectMap4k, DirectMap2M, DirectMap1G**
Breakdown of page table sizes used in the kernel's identity mapping of RAM

### vmallocinfo

Provides information about vmalloced/vmaped areas. One line per area, containing the virtual address range of the area, size in bytes, caller information of the creator, and optional information depending on the kind of area:

| | |
|---|---|
| pages=nr | number of pages |
| phys=addr | if a physical address was specified |
| ioremap | I/O mapping (ioremap() and friends) |
| vmalloc | vmalloc() area |
| vmap | vmap()ed pages |
| user | VM_USERMAP area |
| vpages | buffer for pages pointers was vmalloced (huge area) |
| N<node>=nr | (Only on NUMA kernels) Number of pages allocated on memory node <node> |

```
> cat /proc/vmallocinfo
0xffffc20000000000-0xffffc20000201000 2101248 alloc_large_system_hash+0x204 ...
/0x2c0 pages=512 vmalloc N0=128 N1=128 N2=128 N3=128
0xffffc20000201000-0xffffc20000302000 1052672 alloc_large_system_hash+0x204 ...
/0x2c0 pages=256 vmalloc N0=64 N1=64 N2=64 N3=64
0xffffc20000302000-0xffffc20000304000    8192 acpi_tb_verify_table+0x21/0x4f...
phys=7fee8000 ioremap
0xffffc20000304000-0xffffc20000307000   12288 acpi_tb_verify_table+0x21/0x4f...
phys=7fee7000 ioremap
0xffffc2000031d000-0xffffc2000031f000    8192 init_vdso_vars+0x112/0x210
0xffffc2000031f000-0xffffc2000032b000   49152 cramfs_uncompress_init+0x2e ...
/0x80 pages=11 vmalloc N0=3 N1=3 N2=2 N3=3
0xffffc2000033a000-0xffffc2000033d000   12288 sys_swapon+0x640/0xac0      ...
pages=2 vmalloc N1=2
0xffffc20000347000-0xffffc2000034c000   20480 xt_alloc_table_info+0xfe ...
/0x130 [x_tables] pages=4 vmalloc N0=4
0xffffffffa0000000-0xffffffffa000f000   61440 sys_init_module+0xc27/0x1d00 ...
pages=14 vmalloc N2=14
0xffffffffa000f000-0xffffffffa0014000   20480 sys_init_module+0xc27/0x1d00 ...
pages=4 vmalloc N1=4
0xffffffffa0014000-0xffffffffa0017000   12288 sys_init_module+0xc27/0x1d00 ...
pages=2 vmalloc N1=2
0xffffffffa0017000-0xffffffffa0022000   45056 sys_init_module+0xc27/0x1d00 ...
pages=10 vmalloc N0=10
```

### softirqs

Provides counts of softirq handlers serviced since boot time, for each CPU.

```
> cat /proc/softirqs
               CPU0       CPU1       CPU2       CPU3
     HI:          0          0          0          0
TIMER:       27166      27120      27097      27034
NET_TX:          0          0          0         17
NET_RX:         42          0          0         39
BLOCK:           0          0        107       1121
TASKLET:         0          0          0        290
SCHED:       27035      26983      26971      26746
HRTIMER:         0          0          0          0
    RCU:      1678       1769       2178       2250
```

## 1.3 Networking info in /proc/net

The subdirectory /proc/net follows the usual pattern. Table 1-8 shows the additional values you get for IP version 6 if you configure the kernel to support this. Table 1-9 lists the files and their meaning.

Table 12: Table 1-8: IPv6 info in /proc/net

| File | Content |
|---|---|
| udp6 | UDP sockets (IPv6) |
| tcp6 | TCP sockets (IPv6) |
| raw6 | Raw device statistics (IPv6) |
| igmp6 | IP multicast addresses, which this host joined (IPv6) |
| if_inet6 | List of IPv6 interface addresses |
| ipv6_route | Kernel routing table for IPv6 |
| rt6_stats | Global IPv6 routing tables statistics |
| sockstat6 | Socket statistics (IPv6) |
| snmp6 | Snmp data (IPv6) |

Table 13: Table 1-9: Network info in /proc/net

| File | Content |
|---|---|
| arp | Kernel ARP table |
| dev | network devices with statistics |
| dev_mcast | the Layer2 multicast groups a device is listening too (interface index, label, number of references, number of bound addresses). |
| dev_stat | network device status |
| ip_fwchains | Firewall chain linkage |
| ip_fwnames | Firewall chain names |
| ip_masq | Directory containing the masquerading tables |
| ip_masquerade | Major masquerading table |
| netstat | Network statistics |
| raw | raw device statistics |
| route | Kernel routing table |
| rpc | Directory containing rpc info |
| rt_cache | Routing cache |
| snmp | SNMP data |
| sockstat | Socket statistics |
| softnet_stat | Per-CPU incoming packets queues statistics of online CPUs |
| tcp | TCP sockets |
| udp | UDP sockets |
| unix | UNIX domain sockets |
| wireless | Wireless interface data (Wavelan etc) |
| igmp | IP multicast addresses, which this host joined |
| psched | Global packet scheduler parameters. |
| netlink | List of PF_NETLINK sockets |
| ip_mr_vifs | List of multicast virtual interfaces |
| ip_mr_cache | List of multicast routing cache |

You can use this information to see which network devices are available in your system and how much traffic was routed over those devices:

```
> cat /proc/net/dev
Inter-|Receive                                                   |[...
 face |bytes    packets errs drop fifo frame compressed multicast|[...
    lo:  908188   5596     0    0    0     0          0         0 [...
  ppp0:15475140  20721   410    0    0   410          0         0 [...
  eth0:  614530   7085     0    0    0     0          0         1 [...

...] Transmit
...] bytes    packets errs drop fifo colls carrier compressed
...]  908188     5596    0    0    0     0       0          0
...] 1375103    17405    0    0    0     0       0          0
...] 1703981     5535    0    0    0     3       0          0
```

In addition, each Channel Bond interface has its own directory. For example, the bond0 device will have a directory called /proc/net/bond0/. It will contain information that is specific to that bond, such as the current slaves of the bond, the link status of the slaves, and how many times the slaves link has failed.

### 1.4 SCSI info

If you have a SCSI or ATA host adapter in your system, you'll find a subdirectory named after the driver for this adapter in /proc/scsi. You'll also see a list of all recognized SCSI devices in /proc/scsi:

```
>cat /proc/scsi/scsi
Attached devices:
Host: scsi0 Channel: 00 Id: 00 Lun: 00
  Vendor: IBM      Model: DGHS09U          Rev: 03E0
  Type:   Direct-Access                    ANSI SCSI revision: 03
Host: scsi0 Channel: 00 Id: 06 Lun: 00
  Vendor: PIONEER  Model: CD-ROM DR-U06S   Rev: 1.04
  Type:   CD-ROM                           ANSI SCSI revision: 02
```

The directory named after the driver has one file for each adapter found in the system. These files contain information about the controller, including the used IRQ and the IO address range. The amount of information shown is dependent on the adapter you use. The example shows the output for an Adaptec AHA-2940 SCSI adapter:

```
> cat /proc/scsi/aic7xxx/0

Adaptec AIC7xxx driver version: 5.1.19/3.2.4
Compile Options:
  TCQ Enabled By Default : Disabled
  AIC7XXX_PROC_STATS     : Disabled
  AIC7XXX_RESET_DELAY    : 5
Adapter Configuration:
           SCSI Adapter: Adaptec AHA-294X Ultra SCSI host adapter
                         Ultra Wide Controller
    PCI MMAPed I/O Base: 0xeb001000
 Adapter SEEPROM Config: SEEPROM found and used.
      Adaptec SCSI BIOS: Enabled
                    IRQ: 10
                   SCBs: Active 0, Max Active 2,
                         Allocated 15, HW 16, Page 255
             Interrupts: 160328
      BIOS Control Word: 0x18b6
   Adapter Control Word: 0x005b
   Extended Translation: Enabled
Disconnect Enable Flags: 0xffff
     Ultra Enable Flags: 0x0001
 Tag Queue Enable Flags: 0x0000
Ordered Queue Tag Flags: 0x0000
Default Tag Queue Depth: 8
    Tagged Queue By Device array for aic7xxx host instance 0:
      {255,255,255,255,255,255,255,255,255,255,255,255,255,255,255,255}
    Actual queue depth per device for aic7xxx host instance 0:
      {1,1,1,1,1,1,1,1,1,1,1,1,1,1,1,1}
Statistics:
(scsi0:0:0:0)
  Device using Wide/Sync transfers at 40.0 MByte/sec, offset 8
```

```
  Transinfo settings: current(12/8/1/0), goal(12/8/1/0), user(12/15/1/0)
  Total transfers 160151 (74577 reads and 85574 writes)
(scsi0:0:6:0)
  Device using Narrow/Sync transfers at 5.0 MByte/sec, offset 15
  Transinfo settings: current(50/15/0/0), goal(50/15/0/0), user(50/15/0/0)
  Total transfers 0 (0 reads and 0 writes)
```

### 1.5 Parallel port info in /proc/parport

The directory /proc/parport contains information about the parallel ports of your system. It has one subdirectory for each port, named after the port number (0,1,2,...).

These directories contain the four files shown in Table 1-10.

Table 14: Table 1-10: Files in /proc/parport

| File | Content |
|---|---|
| autoprobe | Any IEEE-1284 device ID information that has been acquired. |
| devices | list of the device drivers using that port. A + will appear by the name of the device currently using the port (it might not appear against any). |
| hardware | Parallel port's base address, IRQ line and DMA channel. |
| irq | IRQ that parport is using for that port. This is in a separate file to allow you to alter it by writing a new value in (IRQ number or none). |

### 1.6 TTY info in /proc/tty

Information about the available and actually used tty's can be found in the directory /proc/tty. You'll find entries for drivers and line disciplines in this directory, as shown in Table 1-11.

Table 15: Table 1-11: Files in /proc/tty

| File | Content |
|---|---|
| drivers | list of drivers and their usage |
| ldiscs | registered line disciplines |
| driver/serial | usage statistic and status of single tty lines |

To see which tty's are currently in use, you can simply look into the file /proc/tty/drivers:

```
> cat /proc/tty/drivers
pty_slave            /dev/pts      136   0-255 pty:slave
pty_master           /dev/ptm      128   0-255 pty:master
pty_slave            /dev/ttyp       3   0-255 pty:slave
pty_master           /dev/pty        2   0-255 pty:master
serial               /dev/cua        5   64-67 serial:callout
serial               /dev/ttyS       4   64-67 serial
/dev/tty0            /dev/tty0       4       0 system:vtmaster
/dev/ptmx            /dev/ptmx       5       2 system
/dev/console         /dev/console    5       1 system:console
/dev/tty             /dev/tty        5       0 system:/dev/tty
unknown              /dev/tty        4    1-63 console
```

### 1.7 Miscellaneous kernel statistics in /proc/stat

Various pieces of information about kernel activity are available in the /proc/stat file. All of the numbers reported in this file are aggregates since the system first booted. For a quick look, simply cat the file:

```
> cat /proc/stat
cpu  237902850 368826709 106375398 1873517540 1135548 0 14507935 0 0 0
cpu0 60045249 91891769 26331539 468411416 495718 0 5739640 0 0 0
cpu1 59746288 91759249 26609887 468860630 312281 0 4384817 0 0 0
cpu2 59489247 92985423 26904446 467808813 171668 0 2268998 0 0 0
cpu3 58622065 92190267 26529524 468436680 155879 0 2114478 0 0 0
intr 8688370575 8 3373 0 0 0 0 0 0 1 40791 0 0 353317 0 0 0 0 224789828 0 0 0␣
↪0 0 0 0 0 0 0 0 0 0 0 0 0 0 0 0 0 0 0 0 0 0 0 0 0 0 0 0 0 0 0 0 0 0 0 0 0␣
↪0 0 0 0 0 0 0 0 0 0 0 0 0 0 0 0 0 0 0 0 0 0 0 0 0 0 0 0 0 0 0 0 0 0 0 0 0␣
↪0 0 0 0 0 0 0 0 0 0 0 0 0 0 0 0 0 0 0 0 0 0 0 0 0 0 0 0 0 190974333 41958554␣
↪123983334 43 0 224593 0 0 0 <more 0's deleted>
ctxt 22848221062
btime 1605316999
processes 746787147
procs_running 2
procs_blocked 0
softirq 12121874454 100099120 3938138295 127375644 2795979 187870761 0␣
↪173808342 3072582055 52608 224184354
```

The very first "cpu" line aggregates the numbers in all of the other "cpuN" lines. These numbers identify the amount of time the CPU has spent performing different kinds of work. Time units are in USER_HZ (typically hundredths of a second). The meanings of the columns are as follows, from left to right:

- user: normal processes executing in user mode
- nice: niced processes executing in user mode
- system: processes executing in kernel mode
- idle: twiddling thumbs
- iowait: In a word, iowait stands for waiting for I/O to complete. But there are several problems:

  1. CPU will not wait for I/O to complete, iowait is the time that a task is waiting for I/O to complete. When CPU goes into idle state for outstanding task I/O, another task will be scheduled on this CPU.
  2. In a multi-core CPU, the task waiting for I/O to complete is not running on any CPU, so the iowait of each CPU is difficult to calculate.
  3. The value of iowait field in /proc/stat will decrease in certain conditions.

  So, the iowait is not reliable by reading from /proc/stat.
- irq: servicing interrupts
- softirq: servicing softirqs
- steal: involuntary wait

- guest: running a normal guest
- guest_nice: running a niced guest

The ”intr” line gives counts of interrupts serviced since boot time, for each of the possible system interrupts. The first column is the total of all interrupts serviced including unnumbered architecture specific interrupts; each subsequent column is the total for that particular numbered interrupt. Unnumbered interrupts are not shown, only summed into the total.

The ”ctxt” line gives the total number of context switches across all CPUs.

The ”btime” line gives the time at which the system booted, in seconds since the Unix epoch.

The ”processes” line gives the number of processes and threads created, which includes (but is not limited to) those created by calls to the fork() and clone() system calls.

The ”procs_running” line gives the total number of threads that are running or ready to run (i.e., the total number of runnable threads).

The ”procs_blocked” line gives the number of processes currently blocked, waiting for I/O to complete.

The ”softirq” line gives counts of softirqs serviced since boot time, for each of the possible system softirqs. The first column is the total of all softirqs serviced; each subsequent column is the total for that particular softirq.

### 1.8 Ext4 file system parameters

Information about mounted ext4 file systems can be found in /proc/fs/ext4. Each mounted filesystem will have a directory in /proc/fs/ext4 based on its device name (i.e., /proc/fs/ext4/hdc or /proc/fs/ext4/sda9 or /proc/fs/ext4/dm-0). The files in each per-device directory are shown in Table 1-12, below.

Table 16: Table 1-12: Files in /proc/fs/ext4/<devname>

| File | Content |
|---|---|
| mb_groups | details of multiblock allocator buddy cache of free blocks |

### 1.9 /proc/consoles

Shows registered system console lines.

To see which character device lines are currently used for the system console /dev/console, you may simply look into the file /proc/consoles:

```
> cat /proc/consoles
tty0                 -WU (ECp)       4:7
ttyS0                -W- (Ep)        4:64
```

The columns are:

| device | name of the device |
|---|---|
| operations | • R = can do read operations<br>• W = can do write operations<br>• U = can do unblank |
| flags | • E = it is enabled<br>• C = it is preferred console<br>• B = it is primary boot console<br>• p = it is used for printk buffer<br>• b = it is not a TTY but a Braille device<br>• a = it is safe to use when cpu is offline |
| major:minor | major and minor number of the device separated by a colon |

#### Summary

The /proc file system serves information about the running system. It not only allows access to process data but also allows you to request the kernel status by reading files in the hierarchy.

The directory structure of /proc reflects the types of information and makes it easy, if not obvious, where to look for specific data.

### 3.43.3 Chapter 2: Modifying System Parameters

#### In This Chapter

- Modifying kernel parameters by writing into files found in /proc/sys
- Exploring the files which modify certain parameters
- Review of the /proc/sys file tree

---

A very interesting part of /proc is the directory /proc/sys. This is not only a source of information, it also allows you to change parameters within the kernel. Be very careful when attempting this. You can optimize your system, but you can also cause it to crash. Never alter kernel parameters on a production system. Set up a development machine and test to make sure that everything works the way you want it to. You may have no alternative but to reboot the machine once an error has been made.

To change a value, simply echo the new value into the file. You need to be root to do this. You can create your own boot script to perform this every time your system boots.

The files in /proc/sys can be used to fine tune and monitor miscellaneous and general things in the operation of the Linux kernel. Since some of the files can inadvertently disrupt your system, it is advisable to read both documentation and source before actually making adjustments. In any case, be very careful when writing to any of these files. The entries in /proc may change slightly between the 2.1.* and the 2.2 kernel, so if there is any doubt review the kernel documentation in the directory linux/Documentation. This chapter is heavily based on

the documentation included in the pre 2.2 kernels, and became part of it in version 2.2.1 of the Linux kernel.

Please see: Documentation/admin-guide/sysctl/ directory for descriptions of these entries.

#### Summary

Certain aspects of kernel behavior can be modified at runtime, without the need to recompile the kernel, or even to reboot the system. The files in the /proc/sys tree can not only be read, but also modified. You can use the echo command to write value into these files, thereby changing the default settings of the kernel.

### 3.43.4 Chapter 3: Per-process Parameters

#### 3.1 /proc/<pid>/oom_adj & /proc/<pid>/oom_score_adj- Adjust the oom-killer score

These files can be used to adjust the badness heuristic used to select which process gets killed in out of memory (oom) conditions.

The badness heuristic assigns a value to each candidate task ranging from 0 (never kill) to 1000 (always kill) to determine which process is targeted. The units are roughly a proportion along that range of allowed memory the process may allocate from based on an estimation of its current memory and swap use. For example, if a task is using all allowed memory, its badness score will be 1000. If it is using half of its allowed memory, its score will be 500.

The amount of "allowed" memory depends on the context in which the oom killer was called. If it is due to the memory assigned to the allocating task's cpuset being exhausted, the allowed memory represents the set of mems assigned to that cpuset. If it is due to a mempolicy's node(s) being exhausted, the allowed memory represents the set of mempolicy nodes. If it is due to a memory limit (or swap limit) being reached, the allowed memory is that configured limit. Finally, if it is due to the entire system being out of memory, the allowed memory represents all allocatable resources.

The value of /proc/<pid>/oom_score_adj is added to the badness score before it is used to determine which task to kill. Acceptable values range from -1000 (OOM_SCORE_ADJ_MIN) to +1000 (OOM_SCORE_ADJ_MAX). This allows userspace to polarize the preference for oom killing either by always preferring a certain task or completely disabling it. The lowest possible value, -1000, is equivalent to disabling oom killing entirely for that task since it will always report a badness score of 0.

Consequently, it is very simple for userspace to define the amount of memory to consider for each task. Setting a /proc/<pid>/oom_score_adj value of +500, for example, is roughly equivalent to allowing the remainder of tasks sharing the same system, cpuset, mempolicy, or memory controller resources to use at least 50% more memory. A value of -500, on the other hand, would be roughly equivalent to discounting 50% of the task's allowed memory from being considered as scoring against the task.

For backwards compatibility with previous kernels, /proc/<pid>/oom_adj may also be used to tune the badness score. Its acceptable values range from -16 (OOM_ADJUST_MIN) to +15 (OOM_ADJUST_MAX) and a special value of -17 (OOM_DISABLE) to disable oom killing entirely for that task. Its value is scaled linearly with /proc/<pid>/oom_score_adj.

The value of /proc/<pid>/oom_score_adj may be reduced no lower than the last value set by a CAP_SYS_RESOURCE process. To reduce the value any lower requires CAP_SYS_RESOURCE.

### 3.2 /proc/<pid>/oom_score - Display current oom-killer score

This file can be used to check the current score used by the oom-killer for any given <pid>. Use it together with /proc/<pid>/oom_score_adj to tune which process should be killed in an out-of-memory situation.

Please note that the exported value includes oom_score_adj so it is effectively in range [0,2000].

### 3.3 /proc/<pid>/io - Display the IO accounting fields

This file contains IO statistics for each running process.

#### Example

```
test:/tmp # dd if=/dev/zero of=/tmp/test.dat &
[1] 3828

test:/tmp # cat /proc/3828/io
rchar: 323934931
wchar: 323929600
syscr: 632687
syscw: 632675
read_bytes: 0
write_bytes: 323932160
cancelled_write_bytes: 0
```

#### Description

#### rchar

I/O counter: chars read The number of bytes which this task has caused to be read from storage. This is simply the sum of bytes which this process passed to read() and pread(). It includes things like tty IO and it is unaffected by whether or not actual physical disk IO was required (the read might have been satisfied from pagecache).

### wchar

I/O counter: chars written The number of bytes which this task has caused, or shall cause to be written to disk. Similar caveats apply here as with rchar.

### syscr

I/O counter: read syscalls Attempt to count the number of read I/O operations, i.e. syscalls like read() and pread().

### syscw

I/O counter: write syscalls Attempt to count the number of write I/O operations, i.e. syscalls like write() and pwrite().

### read_bytes

I/O counter: bytes read Attempt to count the number of bytes which this process really did cause to be fetched from the storage layer. Done at the submit_bio() level, so it is accurate for block-backed filesystems. <please add status regarding NFS and CIFS at a later time>

### write_bytes

I/O counter: bytes written Attempt to count the number of bytes which this process caused to be sent to the storage layer. This is done at page-dirtying time.

### cancelled_write_bytes

The big inaccuracy here is truncate. If a process writes 1MB to a file and then deletes the file, it will in fact perform no writeout. But it will have been accounted as having caused 1MB of write. In other words: The number of bytes which this process caused to not happen, by truncating pagecache. A task can cause ”negative” IO too. If this task truncates some dirty pagecache, some IO which another task has been accounted for (in its write_bytes) will not be happening. We _could_ just subtract that from the truncating task's write_bytes, but there is information loss in doing that.

---

**Note:** At its current implementation state, this is a bit racy on 32-bit machines: if process A reads process B's /proc/pid/io while process B is updating one of those 64-bit counters, process A could see an intermediate result.

---

More information about this can be found within the taskstats documentation in Documentation/accounting.

### 3.4 /proc/<pid>/coredump_filter - Core dump filtering settings

When a process is dumped, all anonymous memory is written to a core file as long as the size of the core file isn't limited. But sometimes we don't want to dump some memory segments, for example, huge shared memory or DAX. Conversely, sometimes we want to save file-backed memory segments into a core file, not only the individual files.

/proc/<pid>/coredump_filter allows you to customize which memory segments will be dumped when the <pid> process is dumped. coredump_filter is a bitmask of memory types. If a bit of the bitmask is set, memory segments of the corresponding memory type are dumped, otherwise they are not dumped.

The following 9 memory types are supported:

- (bit 0) anonymous private memory
- (bit 1) anonymous shared memory
- (bit 2) file-backed private memory
- (bit 3) file-backed shared memory
- (bit 4) ELF header pages in file-backed private memory areas (it is effective only if the bit 2 is cleared)
- (bit 5) hugetlb private memory
- (bit 6) hugetlb shared memory
- (bit 7) DAX private memory
- (bit 8) DAX shared memory

Note that MMIO pages such as frame buffer are never dumped and vDSO pages are always dumped regardless of the bitmask status.

Note that bits 0-4 don't affect hugetlb or DAX memory. hugetlb memory is only affected by bit 5-6, and DAX is only affected by bits 7-8.

The default value of coredump_filter is 0x33; this means all anonymous memory segments, ELF header pages and hugetlb private memory are dumped.

If you don't want to dump all shared memory segments attached to pid 1234, write 0x31 to the process's proc file:

```
$ echo 0x31 > /proc/1234/coredump_filter
```

When a new process is created, the process inherits the bitmask status from its parent. It is useful to set up coredump_filter before the program runs. For example:

```
$ echo 0x7 > /proc/self/coredump_filter
$ ./some_program
```

### 3.5 /proc/<pid>/mountinfo - Information about mounts

This file contains lines of the form:

```
36 35 98:0 /mnt1 /mnt2 rw,noatime master:1 - ext3 /dev/root rw,errors=continue
(1)(2)(3)   (4)   (5)      (6)     (n…m) (m+1)(m+2) (m+3)         (m+4)

(1)   mount ID:        unique identifier of the mount (may be reused after␣
↪umount)
(2)   parent ID:       ID of parent (or of self for the top of the mount tree)
(3)   major:minor:     value of st_dev for files on filesystem
(4)   root:            root of the mount within the filesystem
(5)   mount point:     mount point relative to the process's root
(6)   mount options:   per mount options
(n…m) optional fields: zero or more fields of the form "tag[:value]"
(m+1) separator:       marks the end of the optional fields
(m+2) filesystem type: name of filesystem of the form "type[.subtype]"
(m+3) mount source:    filesystem specific information or "none"
(m+4) super options:   per super block options
```

Parsers should ignore all unrecognised optional fields. Currently the possible optional fields are:

| | |
|---|---|
| shared:X | mount is shared in peer group X |
| master:X | mount is slave to peer group X |
| propagate_from:X | mount is slave and receives propagation from peer group X[1] |
| unbindable | mount is unbindable |

For more information on mount propagation see:

*Shared Subtrees*

[1] X is the closest dominant peer group under the process's root. If X is the immediate master of the mount, or if there's no dominant peer group under the same root, then only the ”master:X” field is present and not the ”propagate_from:X” field.

### 3.6 /proc/<pid>/comm & /proc/<pid>/task/<tid>/comm

These files provide a method to access a task's comm value. It also allows for a task to set its own or one of its thread siblings comm value. The comm value is limited in size compared to the cmdline value, so writing anything longer then the kernel's TASK_COMM_LEN (currently 16 chars) will result in a truncated comm value.

### 3.7 /proc/<pid>/task/<tid>/children - Information about task children

This file provides a fast way to retrieve first level children pids of a task pointed by <pid>/<tid> pair. The format is a space separated stream of pids.

Note the ”first level” here -- if a child has its own children they will not be listed here; one needs to read /proc/<children-pid>/task/<tid>/children to obtain the descendants.

Since this interface is intended to be fast and cheap it doesn't guarantee to provide precise results and some children might be skipped, especially if they've exited right after we printed their pids, so one needs to either stop or freeze processes being inspected if precise results are needed.

### 3.8 /proc/<pid>/fdinfo/<fd> - Information about opened file

This file provides information associated with an opened file. The regular files have at least four fields -- 'pos', 'flags', 'mnt_id' and 'ino'. The 'pos' represents the current offset of the opened file in decimal form [see lseek(2) for details], 'flags' denotes the octal O_xxx mask the file has been created with [see open(2) for details] and 'mnt_id' represents mount ID of the file system containing the opened file [see 3.5 /proc/<pid>/mountinfo for details]. 'ino' represents the inode number of the file.

A typical output is:

```
pos:    0
flags:  0100002
mnt_id: 19
ino:    63107
```

All locks associated with a file descriptor are shown in its fdinfo too:

```
lock:       1: FLOCK  ADVISORY  WRITE 359 00:13:11691 0 EOF
```

The files such as eventfd, fsnotify, signalfd, epoll among the regular pos/flags pair provide additional information particular to the objects they represent.

### Eventfd files

```
pos:    0
flags:  04002
mnt_id: 9
ino:    63107
eventfd-count:  5a
```

where 'eventfd-count' is hex value of a counter.

### Signalfd files

```
pos:    0
flags:  04002
mnt_id: 9
ino:    63107
sigmask:        0000000000000200
```

where 'sigmask' is hex value of the signal mask associated with a file.

### Epoll files

```
pos:    0
flags:  02
mnt_id: 9
ino:    63107
tfd:        5 events:       1d data: ffffffffffffffff pos:0 ino:61af sdev:7
```

where 'tfd' is a target file descriptor number in decimal form, 'events' is events mask being watched and the 'data' is data associated with a target [see epoll(7) for more details].

The 'pos' is current offset of the target file in decimal form [see lseek(2)], 'ino' and 'sdev' are inode and device numbers where target file resides, all in hex format.

### Fsnotify files

For inotify files the format is the following:

```
pos:    0
flags:  02000000
mnt_id: 9
ino:    63107
inotify wd:3 ino:9e7e sdev:800013 mask:800afce ignored_mask:0 fhandle-bytes:8␣
↪fhandle-type:1 f_handle:7e9e0000640d1b6d
```

where 'wd' is a watch descriptor in decimal form, i.e. a target file descriptor number, 'ino' and 'sdev' are inode and device where the target file resides and the 'mask' is the mask of events, all in hex form [see inotify(7) for more details].

If the kernel was built with exportfs support, the path to the target file is encoded as a file handle. The file handle is provided by three fields 'fhandle-bytes', 'fhandle-type' and 'f_handle', all in hex format.

If the kernel is built without exportfs support the file handle won't be printed out.

If there is no inotify mark attached yet the 'inotify' line will be omitted.

For fanotify files the format is:

```
pos:    0
flags:  02
mnt_id: 9
ino:    63107
fanotify flags:10 event-flags:0
fanotify mnt_id:12 mflags:40 mask:38 ignored_mask:40000003
fanotify ino:4f969 sdev:800013 mflags:0 mask:3b ignored_mask:40000000 fhandle-
 ↪bytes:8 fhandle-type:1 f_handle:69f90400c275b5b4
```

where fanotify 'flags' and 'event-flags' are values used in fanotify_init call, 'mnt_id' is the mount point identifier, 'mflags' is the value of flags associated with mark which are tracked separately from events mask. 'ino' and 'sdev' are target inode and device, 'mask' is the events mask and 'ignored_mask' is the mask of events which are to be ignored. All are in hex format. Incorporation of 'mflags', 'mask' and 'ignored_mask' provide information about flags and mask used in fanotify_mark call [see fsnotify manpage for details].

While the first three lines are mandatory and always printed, the rest is optional and may be omitted if no marks created yet.

### Timerfd files

```
pos:    0
flags:  02
mnt_id: 9
ino:    63107
clockid: 0
ticks: 0
settime flags: 01
it_value: (0, 49406829)
it_interval: (1, 0)
```

where 'clockid' is the clock type and 'ticks' is the number of the timer expirations that have occurred [see timerfd_create(2) for details]. 'settime flags' are flags in octal form been used to setup the timer [see timerfd_settime(2) for details]. 'it_value' is remaining time until the timer expiration. 'it_interval' is the interval for the timer. Note the timer might be set up with TIMER_ABSTIME option which will be shown in 'settime flags', but 'it_value' still exhibits timer's remaining time.

#### DMA Buffer files

```
pos:    0
flags:  04002
mnt_id: 9
ino:    63107
size:   32768
count:  2
exp_name:  system-heap
```

where 'size' is the size of the DMA buffer in bytes. 'count' is the file count of the DMA buffer file. 'exp_name' is the name of the DMA buffer exporter.

### 3.9 /proc/<pid>/map_files - Information about memory mapped files

This directory contains symbolic links which represent memory mapped files the process is maintaining. Example output:

```
| lr-------- 1 root root 64 Jan 27 11:24 333c600000-333c620000 -> /usr/lib64/
 ↪ld-2.18.so
| lr-------- 1 root root 64 Jan 27 11:24 333c81f000-333c820000 -> /usr/lib64/
 ↪ld-2.18.so
| lr-------- 1 root root 64 Jan 27 11:24 333c820000-333c821000 -> /usr/lib64/
 ↪ld-2.18.so
| ...
| lr-------- 1 root root 64 Jan 27 11:24 35d0421000-35d0422000 -> /usr/lib64/
 ↪libselinux.so.1
| lr-------- 1 root root 64 Jan 27 11:24 400000-41a000 -> /usr/bin/ls
```

The name of a link represents the virtual memory bounds of a mapping, i.e. vm_area_struct::vm_start-vm_area_struct::vm_end.

The main purpose of the map_files is to retrieve a set of memory mapped files in a fast way instead of parsing /proc/<pid>/maps or /proc/<pid>/smaps, both of which contain many more records. At the same time one can open(2) mappings from the listings of two processes and comparing their inode numbers to figure out which anonymous memory areas are actually shared.

### 3.10 /proc/<pid>/timerslack_ns - Task timerslack value

This file provides the value of the task's timerslack value in nanoseconds. This value specifies an amount of time that normal timers may be deferred in order to coalesce timers and avoid unnecessary wakeups.

This allows a task's interactivity vs power consumption tradeoff to be adjusted.

Writing 0 to the file will set the task's timerslack to the default value.

Valid values are from 0 - ULLONG_MAX

An application setting the value must have PTRACE_MODE_ATTACH_FSCREDS level permissions on the task specified to change its timerslack_ns value.

### 3.11 /proc/<pid>/patch_state - Livepatch patch operation state

When CONFIG_LIVEPATCH is enabled, this file displays the value of the patch state for the task.

A value of '-1' indicates that no patch is in transition.

A value of '0' indicates that a patch is in transition and the task is unpatched. If the patch is being enabled, then the task hasn't been patched yet. If the patch is being disabled, then the task has already been unpatched.

A value of '1' indicates that a patch is in transition and the task is patched. If the patch is being enabled, then the task has already been patched. If the patch is being disabled, then the task hasn't been unpatched yet.

### 3.12 /proc/<pid>/arch_status - task architecture specific status

When CONFIG_PROC_PID_ARCH_STATUS is enabled, this file displays the architecture specific status of the task.

#### Example

```
$ cat /proc/6753/arch_status
AVX512_elapsed_ms:      8
```

#### Description

#### x86 specific entries

#### AVX512_elapsed_ms

If AVX512 is supported on the machine, this entry shows the milliseconds elapsed since the last time AVX512 usage was recorded. The recording happens on a best effort basis when a task is scheduled out. This means that the value depends on two factors:

1) The time which the task spent on the CPU without being scheduled out. With CPU isolation and a single runnable task this can take several seconds.

2) The time since the task was scheduled out last. Depending on the reason for being scheduled out (time slice exhausted, syscall ...) this can be arbitrary long time.

As a consequence the value cannot be considered precise and authoritative information. The application which uses this information has to be aware of the overall scenario on the system in order to determine whether a task is a real AVX512 user or not. Precise information can be obtained with performance counters.

A special value of '-1' indicates that no AVX512 usage was recorded, thus the task is unlikely an AVX512 user, but depends on the workload and the scheduling scenario, it also could be a false negative mentioned above.

### 3.13 /proc/<pid>/fd - List of symlinks to open files

This directory contains symbolic links which represent open files the process is maintaining. Example output:

```
lr-x------ 1 root root 64 Sep 20 17:53 0 -> /dev/null
l-wx------ 1 root root 64 Sep 20 17:53 1 -> /dev/null
lrwx------ 1 root root 64 Sep 20 17:53 10 -> 'socket:[12539]'
lrwx------ 1 root root 64 Sep 20 17:53 11 -> 'socket:[12540]'
lrwx------ 1 root root 64 Sep 20 17:53 12 -> 'socket:[12542]'
```

The number of open files for the process is stored in 'size' member of stat() output for /proc/<pid>/fd for fast access. ----------

## 3.43.5 Chapter 4: Configuring procfs

### 4.1 Mount options

The following mount options are supported:

| | |
|---|---|
| hidepid= | Set /proc/<pid>/ access mode. |
| gid= | Set the group authorized to learn processes information. |
| subset= | Show only the specified subset of procfs. |

hidepid=off or hidepid=0 means classic mode - everybody may access all /proc/<pid>/ directories (default).

hidepid=noaccess or hidepid=1 means users may not access any /proc/<pid>/ directories but their own. Sensitive files like cmdline, sched*, status are now protected against other users. This makes it impossible to learn whether any user runs specific program (given the program doesn't reveal itself by its behaviour). As an additional bonus, as /proc/<pid>/cmdline is unaccessible for other users, poorly written programs passing sensitive information via program arguments are now protected against local eavesdroppers.

hidepid=invisible or hidepid=2 means hidepid=1 plus all /proc/<pid>/ will be fully invisible to other users. It doesn't mean that it hides a fact whether a process with a specific pid value exists (it can be learned by other means, e.g. by "kill -0 $PID"), but it hides process' uid and gid, which may be learned by stat()'ing /proc/<pid>/ otherwise. It greatly complicates an intruder's task of gathering information about running processes, whether some daemon runs with elevated privileges, whether other user runs some sensitive program, whether other users run any program at all, etc.

hidepid=ptraceable or hidepid=4 means that procfs should only contain /proc/<pid>/ directories that the caller can ptrace.

gid= defines a group authorized to learn processes information otherwise prohibited by hidepid=. If you use some daemon like identd which needs to learn information about processes information, just add identd to this group.

subset=pid hides all top level files and directories in the procfs that are not related to tasks.

### 3.43.6 Chapter 5: Filesystem behavior

Originally, before the advent of pid namespace, procfs was a global file system. It means that there was only one procfs instance in the system.

When pid namespace was added, a separate procfs instance was mounted in each pid namespace. So, procfs mount options are global among all mountpoints within the same namespace:

```
# grep ^proc /proc/mounts
proc /proc proc rw,relatime,hidepid=2 0 0

# strace -e mount mount -o hidepid=1 -t proc proc /tmp/proc
mount("proc", "/tmp/proc", "proc", 0, "hidepid=1") = 0
+++ exited with 0 +++

# grep ^proc /proc/mounts
proc /proc proc rw,relatime,hidepid=2 0 0
proc /tmp/proc proc rw,relatime,hidepid=2 0 0
```

and only after remounting procfs mount options will change at all mountpoints:

```
# mount -o remount,hidepid=1 -t proc proc /tmp/proc

# grep ^proc /proc/mounts
proc /proc proc rw,relatime,hidepid=1 0 0
proc /tmp/proc proc rw,relatime,hidepid=1 0 0
```

This behavior is different from the behavior of other filesystems.

The new procfs behavior is more like other filesystems. Each procfs mount creates a new procfs instance. Mount options affect own procfs instance. It means that it became possible to have several procfs instances displaying tasks with different filtering options in one pid namespace:

```
# mount -o hidepid=invisible -t proc proc /proc
# mount -o hidepid=noaccess -t proc proc /tmp/proc
# grep ^proc /proc/mounts
proc /proc proc rw,relatime,hidepid=invisible 0 0
proc /tmp/proc proc rw,relatime,hidepid=noaccess 0 0
```

## 3.44 The QNX6 Filesystem

The qnx6fs is used by newer QNX operating system versions. (e.g. Neutrino) It got introduced in QNX 6.4.0 and is used default since 6.4.1.

### 3.44.1 Option

mmi_fs Mount filesystem as used for example by Audi MMI 3G system

### 3.44.2 Specification

qnx6fs shares many properties with traditional Unix filesystems. It has the concepts of blocks, inodes and directories.

On QNX it is possible to create little endian and big endian qnx6 filesystems. This feature makes it possible to create and use a different endianness fs for the target (QNX is used on quite a range of embedded systems) platform running on a different endianness.

The Linux driver handles endianness transparently. (LE and BE)

#### Blocks

The space in the device or file is split up into blocks. These are a fixed size of 512, 1024, 2048 or 4096, which is decided when the filesystem is created.

Blockpointers are 32bit, so the maximum space that can be addressed is 2^32 * 4096 bytes or 16TB

#### The superblocks

The superblock contains all global information about the filesystem. Each qnx6fs got two superblocks, each one having a 64bit serial number. That serial number is used to identify the ”active” superblock. In write mode with reach new snapshot (after each synchronous write), the serial of the new master superblock is increased (old superblock serial + 1)

So basically the snapshot functionality is realized by an atomic final update of the serial number. Before updating that serial, all modifications are done by copying all modified blocks during that specific write request (or period) and building up a new (stable) filesystem structure under the inactive superblock.

Each superblock holds a set of root inodes for the different filesystem parts. (Inode, Bitmap and Longfilenames) Each of these root nodes holds information like total size of the stored data and the addressing levels in that specific tree. If the level value is 0, up to 16 direct blocks can be addressed by each node.

Level 1 adds an additional indirect addressing level where each indirect addressing block holds up to blocksize / 4 bytes pointers to data blocks. Level 2 adds an additional indirect addressing block level (so, already up to 16 * 256 * 256 = 1048576 blocks that can be addressed by such a tree).

Unused block pointers are always set to ~0 - regardless of root node, indirect addressing blocks or inodes.

Data leaves are always on the lowest level. So no data is stored on upper tree levels.

The first Superblock is located at 0x2000. (0x2000 is the bootblock size) The Audi MMI 3G first superblock directly starts at byte 0.

Second superblock position can either be calculated from the superblock information (total number of filesystem blocks) or by taking the highest device address, zeroing the last 3 bytes and then subtracting 0x1000 from that address.

0x1000 is the size reserved for each superblock - regardless of the blocksize of the filesystem.

### Inodes

Each object in the filesystem is represented by an inode. (index node) The inode structure contains pointers to the filesystem blocks which contain the data held in the object and all of the metadata about an object except its longname. (filenames longer than 27 characters) The metadata about an object includes the permissions, owner, group, flags, size, number of blocks used, access time, change time and modification time.

Object mode field is POSIX format. (which makes things easier)

There are also pointers to the first 16 blocks, if the object data can be addressed with 16 direct blocks.

For more than 16 blocks an indirect addressing in form of another tree is used. (scheme is the same as the one used for the superblock root nodes)

The filesize is stored 64bit. Inode counting starts with 1. (while long filename inodes start with 0)

### Directories

A directory is a filesystem object and has an inode just like a file. It is a specially formatted file containing records which associate each name with an inode number.

'.' inode number points to the directory inode

'..' inode number points to the parent directory inode

Eeach filename record additionally got a filename length field.

One special case are long filenames or subdirectory names.

These got set a filename length field of 0xff in the corresponding directory record plus the longfile inode number also stored in that record.

With that longfilename inode number, the longfilename tree can be walked starting with the superblock longfilename root node pointers.

### Special files

Symbolic links are also filesystem objects with inodes. They got a specific bit in the inode mode field identifying them as symbolic link.

The directory entry file inode pointer points to the target file inode.

Hard links got an inode, a directory entry, but a specific mode bit set, no block pointers and the directory file record pointing to the target file inode.

Character and block special devices do not exist in QNX as those files are handled by the QNX kernel/drivers and created in /dev independent of the underlying filesystem.

### Long filenames

Long filenames are stored in a separate addressing tree. The staring point is the longfilename root node in the active superblock.

Each data block (tree leaves) holds one long filename. That filename is limited to 510 bytes. The first two starting bytes are used as length field for the actual filename.

If that structure shall fit for all allowed blocksizes, it is clear why there is a limit of 510 bytes for the actual filename stored.

### Bitmap

The qnx6fs filesystem allocation bitmap is stored in a tree under bitmap root node in the superblock and each bit in the bitmap represents one filesystem block.

The first block is block 0, which starts 0x1000 after superblock start. So for a normal qnx6fs 0x3000 (bootblock + superblock) is the physical address at which block 0 is located.

Bits at the end of the last bitmap block are set to 1, if the device is smaller than addressing space in the bitmap.

### Bitmap system area

The bitmap itself is divided into three parts.

First the system area, that is split into two halves.

Then userspace.

The requirement for a static, fixed preallocated system area comes from how qnx6fs deals with writes.

Each superblock got its own half of the system area. So superblock #1 always uses blocks from the lower half while superblock #2 just writes to blocks represented by the upper half bitmap system area bits.

Bitmap blocks, Inode blocks and indirect addressing blocks for those two tree structures are treated as system blocks.

The rational behind that is that a write request can work on a new snapshot (system area of the inactive - resp. lower serial numbered superblock) while at the same time there is still a complete stable filesystem structure in the other half of the system area.

When finished with writing (a sync write is completed, the maximum sync leap time or a filesystem sync is requested), serial of the previously inactive superblock atomically is increased and the fs switches over to that - then stable declared - superblock.

For all data outside the system area, blocks are just copied while writing.

## 3.45 Ramfs, rootfs and initramfs

October 17, 2005

**Author**
Rob Landley <rob@landley.net>

### 3.45.1 What is ramfs?

Ramfs is a very simple filesystem that exports Linux's disk caching mechanisms (the page cache and dentry cache) as a dynamically resizable RAM-based filesystem.

Normally all files are cached in memory by Linux. Pages of data read from backing store (usually the block device the filesystem is mounted on) are kept around in case it's needed again, but marked as clean (freeable) in case the Virtual Memory system needs the memory for something else. Similarly, data written to files is marked clean as soon as it has been written to backing store, but kept around for caching purposes until the VM reallocates the memory. A similar mechanism (the dentry cache) greatly speeds up access to directories.

With ramfs, there is no backing store. Files written into ramfs allocate dentries and page cache as usual, but there's nowhere to write them to. This means the pages are never marked clean, so they can't be freed by the VM when it's looking to recycle memory.

The amount of code required to implement ramfs is tiny, because all the work is done by the existing Linux caching infrastructure. Basically, you're mounting the disk cache as a filesystem. Because of this, ramfs is not an optional component removable via menuconfig, since there would be negligible space savings.

### 3.45.2 ramfs and ramdisk:

The older ”ram disk” mechanism created a synthetic block device out of an area of RAM and used it as backing store for a filesystem. This block device was of fixed size, so the filesystem mounted on it was of fixed size. Using a ram disk also required unnecessarily copying memory from the fake block device into the page cache (and copying changes back out), as well as creating and destroying dentries. Plus it needed a filesystem driver (such as ext2) to format and interpret this data.

Compared to ramfs, this wastes memory (and memory bus bandwidth), creates unnecessary work for the CPU, and pollutes the CPU caches. (There are tricks to avoid this copying by playing with the page tables, but they're unpleasantly complicated and turn out to be about as expensive as the copying anyway.) More to the point, all the work ramfs is doing has to happen _anyway_, since all file access goes through the page and dentry caches. The RAM disk is simply unnecessary; ramfs is internally much simpler.

Another reason ramdisks are semi-obsolete is that the introduction of loopback devices offered a more flexible and convenient way to create synthetic block devices, now from files instead of from chunks of memory. See losetup (8) for details.

### 3.45.3 ramfs and tmpfs:

One downside of ramfs is you can keep writing data into it until you fill up all memory, and the VM can't free it because the VM thinks that files should get written to backing store (rather than swap space), but ramfs hasn't got any backing store. Because of this, only root (or a trusted user) should be allowed write access to a ramfs mount.

A ramfs derivative called tmpfs was created to add size limits, and the ability to write the data to swap space. Normal users can be allowed write access to tmpfs mounts. See *Tmpfs* for more information.

### 3.45.4 What is rootfs?

Rootfs is a special instance of ramfs (or tmpfs, if that's enabled), which is always present in 2.6 systems. You can't unmount rootfs for approximately the same reason you can't kill the init process; rather than having special code to check for and handle an empty list, it's smaller and simpler for the kernel to just make sure certain lists can't become empty.

Most systems just mount another filesystem over rootfs and ignore it. The amount of space an empty instance of ramfs takes up is tiny.

If CONFIG_TMPFS is enabled, rootfs will use tmpfs instead of ramfs by default. To force ramfs, add "rootfstype=ramfs" to the kernel command line.

### 3.45.5 What is initramfs?

All 2.6 Linux kernels contain a gzipped "cpio" format archive, which is extracted into rootfs when the kernel boots up. After extracting, the kernel checks to see if rootfs contains a file "init", and if so it executes it as PID 1. If found, this init process is responsible for bringing the system the rest of the way up, including locating and mounting the real root device (if any). If rootfs does not contain an init program after the embedded cpio archive is extracted into it, the kernel will fall through to the older code to locate and mount a root partition, then exec some variant of /sbin/init out of that.

All this differs from the old initrd in several ways:

- The old initrd was always a separate file, while the initramfs archive is linked into the linux kernel image. (The directory `linux-*/usr` is devoted to generating this archive during the build.)
- The old initrd file was a gzipped filesystem image (in some file format, such as ext2, that needed a driver built into the kernel), while the new initramfs archive is a gzipped cpio archive (like tar only simpler, see cpio(1) and Documentation/driver-api/early-userspace/buffer-format.rst). The kernel's cpio extraction code is not only extremely small, it's also __init text and data that can be discarded during the boot process.
- The program run by the old initrd (which was called /initrd, not /init) did some setup and then returned to the kernel, while the init program from initramfs is not expected to return

to the kernel. (If /init needs to hand off control it can overmount / with a new root device and exec another init program. See the switch_root utility, below.)

- When switching another root device, initrd would pivot_root and then umount the ramdisk. But initramfs is rootfs: you can neither pivot_root rootfs, nor unmount it. Instead delete everything out of rootfs to free up the space (find -xdev / -exec rm '{}' ';'), overmount rootfs with the new root (cd /newmount; mount --move . /; chroot .), attach stdin/stdout/stderr to the new /dev/console, and exec the new init.

  Since this is a remarkably persnickety process (and involves deleting commands before you can run them), the klibc package introduced a helper program (utils/run_init.c) to do all this for you. Most other packages (such as busybox) have named this command "switch_root".

### 3.45.6 Populating initramfs:

The 2.6 kernel build process always creates a gzipped cpio format initramfs archive and links it into the resulting kernel binary. By default, this archive is empty (consuming 134 bytes on x86).

The config option CONFIG_INITRAMFS_SOURCE (in General Setup in menuconfig, and living in usr/Kconfig) can be used to specify a source for the initramfs archive, which will automatically be incorporated into the resulting binary. This option can point to an existing gzipped cpio archive, a directory containing files to be archived, or a text file specification such as the following example:

```
dir /dev 755 0 0
nod /dev/console 644 0 0 c 5 1
nod /dev/loop0 644 0 0 b 7 0
dir /bin 755 1000 1000
slink /bin/sh busybox 777 0 0
file /bin/busybox initramfs/busybox 755 0 0
dir /proc 755 0 0
dir /sys 755 0 0
dir /mnt 755 0 0
file /init initramfs/init.sh 755 0 0
```

Run "usr/gen_init_cpio" (after the kernel build) to get a usage message documenting the above file format.

One advantage of the configuration file is that root access is not required to set permissions or create device nodes in the new archive. (Note that those two example "file" entries expect to find files named "init.sh" and "busybox" in a directory called "initramfs", under the linux-2.6.* directory. See Documentation/driver-api/early-userspace/early_userspace_support.rst for more details.)

The kernel does not depend on external cpio tools. If you specify a directory instead of a configuration file, the kernel's build infrastructure creates a configuration file from that directory (usr/Makefile calls usr/gen_initramfs.sh), and proceeds to package up that directory using the config file (by feeding it to usr/gen_init_cpio, which is created from usr/gen_init_cpio.c). The kernel's build-time cpio creation code is entirely self-contained, and the kernel's boot-time extractor is also (obviously) self-contained.

The one thing you might need external cpio utilities installed for is creating or extracting your own preprepared cpio files to feed to the kernel build (instead of a config file or directory).

The following command line can extract a cpio image (either by the above script or by the kernel build) back into its component files:

```
cpio -i -d -H newc -F initramfs_data.cpio --no-absolute-filenames
```

The following shell script can create a prebuilt cpio archive you can use in place of the above config file:

```
#!/bin/sh

# Copyright 2006 Rob Landley <rob@landley.net> and TimeSys Corporation.
# Licensed under GPL version 2

if [ $# -ne 2 ]
then
  echo "usage: mkinitramfs directory imagename.cpio.gz"
  exit 1
fi

if [ -d "$1" ]
then
  echo "creating $2 from $1"
  (cd "$1"; find . | cpio -o -H newc | gzip) > "$2"
else
  echo "First argument must be a directory"
  exit 1
fi
```

---

**Note:** The cpio man page contains some bad advice that will break your initramfs archive if you follow it. It says ”A typical way to generate the list of filenames is with the find command; you should give find the -depth option to minimize problems with permissions on directories that are unwritable or not searchable.” Don't do this when creating initramfs.cpio.gz images, it won't work. The Linux kernel cpio extractor won't create files in a directory that doesn't exist, so the directory entries must go before the files that go in those directories. The above script gets them in the right order.

---

### 3.45.7 External initramfs images:

If the kernel has initrd support enabled, an external cpio.gz archive can also be passed into a 2.6 kernel in place of an initrd. In this case, the kernel will autodetect the type (initramfs, not initrd) and extract the external cpio archive into rootfs before trying to run /init.

This has the memory efficiency advantages of initramfs (no ramdisk block device) but the separate packaging of initrd (which is nice if you have non-GPL code you'd like to run from initramfs, without conflating it with the GPL licensed Linux kernel binary).

It can also be used to supplement the kernel's built-in initramfs image. The files in the external archive will overwrite any conflicting files in the built-in initramfs archive. Some distributors

also prefer to customize a single kernel image with task-specific initramfs images, without recompiling.

### 3.45.8 Contents of initramfs:

An initramfs archive is a complete self-contained root filesystem for Linux. If you don't already understand what shared libraries, devices, and paths you need to get a minimal root filesystem up and running, here are some references:

- https://www.tldp.org/HOWTO/Bootdisk-HOWTO/
- https://www.tldp.org/HOWTO/From-PowerUp-To-Bash-Prompt-HOWTO.html
- http://www.linuxfromscratch.org/lfs/view/stable/

The "klibc" package (https://www.kernel.org/pub/linux/libs/klibc) is designed to be a tiny C library to statically link early userspace code against, along with some related utilities. It is BSD licensed.

I use uClibc (https://www.uclibc.org) and busybox (https://www.busybox.net) myself. These are LGPL and GPL, respectively. (A self-contained initramfs package is planned for the busybox 1.3 release.)

In theory you could use glibc, but that's not well suited for small embedded uses like this. (A "hello world" program statically linked against glibc is over 400k. With uClibc it's 7k. Also note that glibc dlopens libnss to do name lookups, even when otherwise statically linked.)

A good first step is to get initramfs to run a statically linked "hello world" program as init, and test it under an emulator like qemu (www.qemu.org) or User Mode Linux, like so:

```
cat > hello.c << EOF
#include <stdio.h>
#include <unistd.h>

int main(int argc, char *argv[])
{
  printf("Hello world!\n");
  sleep(999999999);
}
EOF
gcc -static hello.c -o init
echo init | cpio -o -H newc | gzip > test.cpio.gz
# Testing external initramfs using the initrd loading mechanism.
qemu -kernel /boot/vmlinuz -initrd test.cpio.gz /dev/zero
```

When debugging a normal root filesystem, it's nice to be able to boot with "init=/bin/sh". The initramfs equivalent is "rdinit=/bin/sh", and it's just as useful.

### 3.45.9 Why cpio rather than tar?

This decision was made back in December, 2001. The discussion started here:

http://www.uwsg.iu.edu/hypermail/linux/kernel/0112.2/1538.html

And spawned a second thread (specifically on tar vs cpio), starting here:

http://www.uwsg.iu.edu/hypermail/linux/kernel/0112.2/1587.html

The quick and dirty summary version (which is no substitute for reading the above threads) is:

1) cpio is a standard. It's decades old (from the AT&T days), and already widely used on Linux (inside RPM, Red Hat's device driver disks). Here's a Linux Journal article about it from 1996:

   http://www.linuxjournal.com/article/1213

   It's not as popular as tar because the traditional cpio command line tools require _truly_hideous_ command line arguments. But that says nothing either way about the archive format, and there are alternative tools, such as:

   http://freecode.com/projects/afio

2) The cpio archive format chosen by the kernel is simpler and cleaner (and thus easier to create and parse) than any of the (literally dozens of) various tar archive formats. The complete initramfs archive format is explained in buffer-format.txt, created in usr/gen_init_cpio.c, and extracted in init/initramfs.c. All three together come to less than 26k total of human-readable text.

3) The GNU project standardizing on tar is approximately as relevant as Windows standardizing on zip. Linux is not part of either, and is free to make its own technical decisions.

4) Since this is a kernel internal format, it could easily have been something brand new. The kernel provides its own tools to create and extract this format anyway. Using an existing standard was preferable, but not essential.

5) Al Viro made the decision (quote: "tar is ugly as hell and not going to be supported on the kernel side"):

   http://www.uwsg.iu.edu/hypermail/linux/kernel/0112.2/1540.html

   explained his reasoning:

   - http://www.uwsg.iu.edu/hypermail/linux/kernel/0112.2/1550.html
   - http://www.uwsg.iu.edu/hypermail/linux/kernel/0112.2/1638.html

   and, most importantly, designed and implemented the initramfs code.

### 3.45.10 Future directions:

Today (2.6.16), initramfs is always compiled in, but not always used. The kernel falls back to legacy boot code that is reached only if initramfs does not contain an /init program. The fallback is legacy code, there to ensure a smooth transition and allowing early boot functionality to gradually move to ”early userspace” (I.E. initramfs).

The move to early userspace is necessary because finding and mounting the real root device is complex. Root partitions can span multiple devices (raid or separate journal). They can be out on the network (requiring dhcp, setting a specific MAC address, logging into a server, etc). They can live on removable media, with dynamically allocated major/minor numbers and persistent naming issues requiring a full udev implementation to sort out. They can be compressed, encrypted, copy-on-write, loopback mounted, strangely partitioned, and so on.

This kind of complexity (which inevitably includes policy) is rightly handled in userspace. Both klibc and busybox/uClibc are working on simple initramfs packages to drop into a kernel build.

The klibc package has now been accepted into Andrew Morton's 2.6.17-mm tree. The kernel's current early boot code (partition detection, etc) will probably be migrated into a default initramfs, automatically created and used by the kernel build.

## 3.46 relay interface (formerly relayfs)

The relay interface provides a means for kernel applications to efficiently log and transfer large quantities of data from the kernel to userspace via user-defined 'relay channels'.

A 'relay channel' is a kernel->user data relay mechanism implemented as a set of per-cpu kernel buffers ('channel buffers'), each represented as a regular file ('relay file') in user space. Kernel clients write into the channel buffers using efficient write functions; these automatically log into the current cpu's channel buffer. User space applications mmap() or read() from the relay files and retrieve the data as it becomes available. The relay files themselves are files created in a host filesystem, e.g. debugfs, and are associated with the channel buffers using the API described below.

The format of the data logged into the channel buffers is completely up to the kernel client; the relay interface does however provide hooks which allow kernel clients to impose some structure on the buffer data. The relay interface doesn't implement any form of data filtering - this also is left to the kernel client. The purpose is to keep things as simple as possible.

This document provides an overview of the relay interface API. The details of the function parameters are documented along with the functions in the relay interface code - please see that for details.

### 3.46.1 Semantics

Each relay channel has one buffer per CPU, each buffer has one or more sub-buffers. Messages are written to the first sub-buffer until it is too full to contain a new message, in which case it is written to the next (if available). Messages are never split across sub-buffers. At this point, userspace can be notified so it empties the first sub-buffer, while the kernel continues writing to the next.

When notified that a sub-buffer is full, the kernel knows how many bytes of it are padding i.e. unused space occurring because a complete message couldn't fit into a sub-buffer. Userspace can use this knowledge to copy only valid data.

After copying it, userspace can notify the kernel that a sub-buffer has been consumed.

A relay channel can operate in a mode where it will overwrite data not yet collected by userspace, and not wait for it to be consumed.

The relay channel itself does not provide for communication of such data between userspace and kernel, allowing the kernel side to remain simple and not impose a single interface on userspace. It does provide a set of examples and a separate helper though, described below.

The read() interface both removes padding and internally consumes the read sub-buffers; thus in cases where read(2) is being used to drain the channel buffers, special-purpose communication between kernel and user isn't necessary for basic operation.

One of the major goals of the relay interface is to provide a low overhead mechanism for conveying kernel data to userspace. While the read() interface is easy to use, it's not as efficient as the mmap() approach; the example code attempts to make the tradeoff between the two approaches as small as possible.

### 3.46.2 klog and relay-apps example code

The relay interface itself is ready to use, but to make things easier, a couple simple utility functions and a set of examples are provided.

The relay-apps example tarball, available on the relay sourceforge site, contains a set of self-contained examples, each consisting of a pair of .c files containing boilerplate code for each of the user and kernel sides of a relay application. When combined these two sets of boilerplate code provide glue to easily stream data to disk, without having to bother with mundane housekeeping chores.

The 'klog debugging functions' patch (klog.patch in the relay-apps tarball) provides a couple of high-level logging functions to the kernel which allow writing formatted text or raw data to a channel, regardless of whether a channel to write into exists or not, or even whether the relay interface is compiled into the kernel or not. These functions allow you to put unconditional 'trace' statements anywhere in the kernel or kernel modules; only when there is a 'klog handler' registered will data actually be logged (see the klog and kleak examples for details).

It is of course possible to use the relay interface from scratch, i.e. without using any of the relay-apps example code or klog, but you'll have to implement communication between userspace and kernel, allowing both to convey the state of buffers (full, empty, amount of padding). The read() interface both removes padding and internally consumes the read sub-buffers; thus in cases where read(2) is being used to drain the channel buffers, special-purpose communication between kernel and user isn't necessary for basic operation. Things such as buffer-full conditions would still need to be communicated via some channel though.

klog and the relay-apps examples can be found in the relay-apps tarball on http://relayfs.sourceforge.net

### 3.46.3 The relay interface user space API

The relay interface implements basic file operations for user space access to relay channel buffer data. Here are the file operations that are available and some comments regarding their behavior:

| | |
|---|---|
| open() | enables user to open an _existing_ channel buffer. |
| mmap() | results in channel buffer being mapped into the caller's memory space. Note that you can't do a partial mmap - you must map the entire file, which is NRBUF * SUBBUFSIZE. |
| read() | read the contents of a channel buffer. The bytes read are 'consumed' by the reader, i.e. they won't be available again to subsequent reads. If the channel is being used in no-overwrite mode (the default), it can be read at any time even if there's an active kernel writer. If the channel is being used in overwrite mode and there are active channel writers, results may be unpredictable - users should make sure that all logging to the channel has ended before using read() with overwrite mode. Sub-buffer padding is automatically removed and will not be seen by the reader. |
| sendfile() | transfer data from a channel buffer to an output file descriptor. Sub-buffer padding is automatically removed and will not be seen by the reader. |
| poll() | POLLIN/POLLRDNORM/POLLERR supported. User applications are notified when sub-buffer boundaries are crossed. |
| close() | decrements the channel buffer's refcount. When the refcount reaches 0, i.e. when no process or kernel client has the buffer open, the channel buffer is freed. |

In order for a user application to make use of relay files, the host filesystem must be mounted. For example:

```
mount -t debugfs debugfs /sys/kernel/debug
```

**Note:** the host filesystem doesn't need to be mounted for kernel clients to create or use channels - it only needs to be mounted when user space applications need access to the buffer data.

### 3.46.4 The relay interface kernel API

Here's a summary of the API the relay interface provides to in-kernel clients:

**TBD(curr. line MT:/API/)**

channel management functions:

```
relay_open(base_filename, parent, subbuf_size, n_subbufs,
           callbacks, private_data)
relay_close(chan)
```

```
relay_flush(chan)
relay_reset(chan)
```

channel management typically called on instigation of userspace:

```
relay_subbufs_consumed(chan, cpu, subbufs_consumed)
```

write functions:

```
relay_write(chan, data, length)
__relay_write(chan, data, length)
relay_reserve(chan, length)
```

callbacks:

```
subbuf_start(buf, subbuf, prev_subbuf, prev_padding)
buf_mapped(buf, filp)
buf_unmapped(buf, filp)
create_buf_file(filename, parent, mode, buf, is_global)
remove_buf_file(dentry)
```

helper functions:

```
relay_buf_full(buf)
subbuf_start_reserve(buf, length)
```

### Creating a channel

relay_open() is used to create a channel, along with its per-cpu channel buffers. Each channel buffer will have an associated file created for it in the host filesystem, which can be and mmapped or read from in user space. The files are named basename0...basenameN-1 where N is the number of online cpus, and by default will be created in the root of the filesystem (if the parent param is NULL). If you want a directory structure to contain your relay files, you should create it using the host filesystem's directory creation function, e.g. *debugfs_create_dir()*, and pass the parent directory to relay_open(). Users are responsible for cleaning up any directory structure they create, when the channel is closed - again the host filesystem's directory removal functions should be used for that, e.g. *debugfs_remove()*.

In order for a channel to be created and the host filesystem's files associated with its channel buffers, the user must provide definitions for two callback functions, create_buf_file() and remove_buf_file(). create_buf_file() is called once for each per-cpu buffer from relay_open() and allows the user to create the file which will be used to represent the corresponding channel buffer. The callback should return the dentry of the file created to represent the channel buffer. remove_buf_file() must also be defined; it's responsible for deleting the file(s) created in create_buf_file() and is called during relay_close().

Here are some typical definitions for these callbacks, in this case using debugfs:

```
/*
* create_buf_file() callback.  Creates relay file in debugfs.
*/
static struct dentry *create_buf_file_handler(const char *filename,
```

```
                                        struct dentry *parent,
                                        umode_t mode,
                                        struct rchan_buf *buf,
                                        int *is_global)
{
        return debugfs_create_file(filename, mode, parent, buf,
                                   &relay_file_operations);
}

/*
* remove_buf_file() callback.  Removes relay file from debugfs.
*/
static int remove_buf_file_handler(struct dentry *dentry)
{
        debugfs_remove(dentry);

        return 0;
}

/*
* relay interface callbacks
*/
static struct rchan_callbacks relay_callbacks =
{
        .create_buf_file = create_buf_file_handler,
        .remove_buf_file = remove_buf_file_handler,
};
```

And an example relay_open() invocation using them:

```
chan = relay_open("cpu", NULL, SUBBUF_SIZE, N_SUBBUFS, &relay_callbacks, NULL);
```

If the create_buf_file() callback fails, or isn't defined, channel creation and thus relay_open() will fail.

The total size of each per-cpu buffer is calculated by multiplying the number of sub-buffers by the sub-buffer size passed into relay_open(). The idea behind sub-buffers is that they're basically an extension of double-buffering to N buffers, and they also allow applications to easily implement random-access-on-buffer-boundary schemes, which can be important for some high-volume applications. The number and size of sub-buffers is completely dependent on the application and even for the same application, different conditions will warrant different values for these parameters at different times. Typically, the right values to use are best decided after some experimentation; in general, though, it's safe to assume that having only 1 sub-buffer is a bad idea - you're guaranteed to either overwrite data or lose events depending on the channel mode being used.

The create_buf_file() implementation can also be defined in such a way as to allow the creation of a single 'global' buffer instead of the default per-cpu set. This can be useful for applications interested mainly in seeing the relative ordering of system-wide events without the need to bother with saving explicit timestamps for the purpose of merging/sorting per-cpu files in a postprocessing step.

To have relay_open() create a global buffer, the create_buf_file() implementation should set the value of the is_global outparam to a non-zero value in addition to creating the file that will be used to represent the single buffer. In the case of a global buffer, create_buf_file() and remove_buf_file() will be called only once. The normal channel-writing functions, e.g. relay_write(), can still be used - writes from any cpu will transparently end up in the global buffer - but since it is a global buffer, callers should make sure they use the proper locking for such a buffer, either by wrapping writes in a spinlock, or by copying a write function from relay.h and creating a local version that internally does the proper locking.

The private_data passed into relay_open() allows clients to associate user-defined data with a channel, and is immediately available (including in create_buf_file()) via chan->private_data or buf->chan->private_data.

### Buffer-only channels

These channels have no files associated and can be created with relay_open(NULL, NULL, ...). Such channels are useful in scenarios such as when doing early tracing in the kernel, before the VFS is up. In these cases, one may open a buffer-only channel and then call relay_late_setup_files() when the kernel is ready to handle files, to expose the buffered data to the userspace.

### Channel 'modes'

relay channels can be used in either of two modes - 'overwrite' or 'no-overwrite'. The mode is entirely determined by the implementation of the subbuf_start() callback, as described below. The default if no subbuf_start() callback is defined is 'no-overwrite' mode. If the default mode suits your needs, and you plan to use the read() interface to retrieve channel data, you can ignore the details of this section, as it pertains mainly to mmap() implementations.

In 'overwrite' mode, also known as 'flight recorder' mode, writes continuously cycle around the buffer and will never fail, but will unconditionally overwrite old data regardless of whether it's actually been consumed. In no-overwrite mode, writes will fail, i.e. data will be lost, if the number of unconsumed sub-buffers equals the total number of sub-buffers in the channel. It should be clear that if there is no consumer or if the consumer can't consume sub-buffers fast enough, data will be lost in either case; the only difference is whether data is lost from the beginning or the end of a buffer.

As explained above, a relay channel is made of up one or more per-cpu channel buffers, each implemented as a circular buffer subdivided into one or more sub-buffers. Messages are written into the current sub-buffer of the channel's current per-cpu buffer via the write functions described below. Whenever a message can't fit into the current sub-buffer, because there's no room left for it, the client is notified via the subbuf_start() callback that a switch to a new sub-buffer is about to occur. The client uses this callback to 1) initialize the next sub-buffer if appropriate 2) finalize the previous sub-buffer if appropriate and 3) return a boolean value indicating whether or not to actually move on to the next sub-buffer.

To implement 'no-overwrite' mode, the userspace client would provide an implementation of the subbuf_start() callback something like the following:

```
static int subbuf_start(struct rchan_buf *buf,
                        void *subbuf,
                        void *prev_subbuf,
```

```
                        unsigned int prev_padding)
{
        if (prev_subbuf)
                *((unsigned *)prev_subbuf) = prev_padding;

        if (relay_buf_full(buf))
                return 0;

        subbuf_start_reserve(buf, sizeof(unsigned int));

        return 1;
}
```

If the current buffer is full, i.e. all sub-buffers remain unconsumed, the callback returns 0 to indicate that the buffer switch should not occur yet, i.e. until the consumer has had a chance to read the current set of ready sub-buffers. For the relay_buf_full() function to make sense, the consumer is responsible for notifying the relay interface when sub-buffers have been consumed via relay_subbufs_consumed(). Any subsequent attempts to write into the buffer will again invoke the subbuf_start() callback with the same parameters; only when the consumer has consumed one or more of the ready sub-buffers will relay_buf_full() return 0, in which case the buffer switch can continue.

The implementation of the subbuf_start() callback for 'overwrite' mode would be very similar:

```
static int subbuf_start(struct rchan_buf *buf,
                        void *subbuf,
                        void *prev_subbuf,
                        size_t prev_padding)
{
        if (prev_subbuf)
                *((unsigned *)prev_subbuf) = prev_padding;

        subbuf_start_reserve(buf, sizeof(unsigned int));

        return 1;
}
```

In this case, the relay_buf_full() check is meaningless and the callback always returns 1, causing the buffer switch to occur unconditionally. It's also meaningless for the client to use the relay_subbufs_consumed() function in this mode, as it's never consulted.

The default subbuf_start() implementation, used if the client doesn't define any callbacks, or doesn't define the subbuf_start() callback, implements the simplest possible 'no-overwrite' mode, i.e. it does nothing but return 0.

Header information can be reserved at the beginning of each sub-buffer by calling the subbuf_start_reserve() helper function from within the subbuf_start() callback. This reserved area can be used to store whatever information the client wants. In the example above, room is reserved in each sub-buffer to store the padding count for that sub-buffer. This is filled in for the previous sub-buffer in the subbuf_start() implementation; the padding value for the previous sub-buffer is passed into the subbuf_start() callback along with a pointer to the previous sub-buffer, since the padding value isn't known until a sub-buffer is filled. The subbuf_start()

callback is also called for the first sub-buffer when the channel is opened, to give the client a chance to reserve space in it. In this case the previous sub-buffer pointer passed into the callback will be NULL, so the client should check the value of the prev_subbuf pointer before writing into the previous sub-buffer.

### Writing to a channel

Kernel clients write data into the current cpu's channel buffer using relay_write() or __relay_write(). relay_write() is the main logging function - it uses local_irqsave() to protect the buffer and should be used if you might be logging from interrupt context. If you know you'll never be logging from interrupt context, you can use __relay_write(), which only disables preemption. These functions don't return a value, so you can't determine whether or not they failed - the assumption is that you wouldn't want to check a return value in the fast logging path anyway, and that they'll always succeed unless the buffer is full and no-overwrite mode is being used, in which case you can detect a failed write in the subbuf_start() callback by calling the relay_buf_full() helper function.

relay_reserve() is used to reserve a slot in a channel buffer which can be written to later. This would typically be used in applications that need to write directly into a channel buffer without having to stage data in a temporary buffer beforehand. Because the actual write may not happen immediately after the slot is reserved, applications using relay_reserve() can keep a count of the number of bytes actually written, either in space reserved in the sub-buffers themselves or as a separate array. See the 'reserve' example in the relay-apps tarball at http://relayfs.sourceforge.net for an example of how this can be done. Because the write is under control of the client and is separated from the reserve, relay_reserve() doesn't protect the buffer at all - it's up to the client to provide the appropriate synchronization when using relay_reserve().

### Closing a channel

The client calls relay_close() when it's finished using the channel. The channel and its associated buffers are destroyed when there are no longer any references to any of the channel buffers. relay_flush() forces a sub-buffer switch on all the channel buffers, and can be used to finalize and process the last sub-buffers before the channel is closed.

### Misc

Some applications may want to keep a channel around and re-use it rather than open and close a new channel for each use. relay_reset() can be used for this purpose - it resets a channel to its initial state without reallocating channel buffer memory or destroying existing mappings. It should however only be called when it's safe to do so, i.e. when the channel isn't currently being written to.

Finally, there are a couple of utility callbacks that can be used for different purposes. buf_mapped() is called whenever a channel buffer is mmapped from user space and buf_unmapped() is called when it's unmapped. The client can use this notification to trigger actions within the kernel application, such as enabling/disabling logging to the channel.

### 3.46.5 Resources

For news, example code, mailing list, etc. see the relay interface homepage:

http://relayfs.sourceforge.net

### 3.46.6 Credits

The ideas and specs for the relay interface came about as a result of discussions on tracing involving the following:

Michel Dagenais <michel.dagenais@polymtl.ca> Richard Moore <richardj_moore@uk.ibm.com> Bob Wisniewski <bob@watson.ibm.com> Karim Yaghmour <karim@opersys.com> Tom Zanussi <zanussi@us.ibm.com>

Also thanks to Hubertus Franke for a lot of useful suggestions and bug reports.

## 3.47 ROMFS - ROM File System

This is a quite dumb, read only filesystem, mainly for initial RAM disks of installation disks. It has grown up by the need of having modules linked at boot time. Using this filesystem, you get a very similar feature, and even the possibility of a small kernel, with a file system which doesn't take up useful memory from the router functions in the basement of your office.

For comparison, both the older minix and xiafs (the latter is now defunct) filesystems, compiled as module need more than 20000 bytes, while romfs is less than a page, about 4000 bytes (assuming i586 code). Under the same conditions, the msdos filesystem would need about 30K (and does not support device nodes or symlinks), while the nfs module with nfsroot is about 57K. Furthermore, as a bit unfair comparison, an actual rescue disk used up 3202 blocks with ext2, while with romfs, it needed 3079 blocks.

To create such a file system, you'll need a user program named genromfs. It is available on http://romfs.sourceforge.net/

As the name suggests, romfs could be also used (space-efficiently) on various read-only media, like (E)EPROM disks if someone will have the motivation.. :)

However, the main purpose of romfs is to have a very small kernel, which has only this filesystem linked in, and then can load any module later, with the current module utilities. It can also be used to run some program to decide if you need SCSI devices, and even IDE or floppy drives can be loaded later if you use the ”initrd”--initial RAM disk--feature of the kernel. This would not be really news flash, but with romfs, you can even spare off your ext2 or minix or maybe even affs filesystem until you really know that you need it.

For example, a distribution boot disk can contain only the cd disk drivers (and possibly the SCSI drivers), and the ISO 9660 filesystem module. The kernel can be small enough, since it doesn't have other filesystems, like the quite large ext2fs module, which can then be loaded off the CD at a later stage of the installation. Another use would be for a recovery disk, when you are reinstalling a workstation from the network, and you will have all the tools/modules available from a nearby server, so you don't want to carry two disks for this purpose, just because it won't fit into ext2.

romfs operates on block devices as you can expect, and the underlying structure is very simple. Every accessible structure begins on 16 byte boundaries for fast access. The minimum space

a file will take is 32 bytes (this is an empty file, with a less than 16 character name). The maximum overhead for any non-empty file is the header, and the 16 byte padding for the name and the contents, also 16+14+15 = 45 bytes. This is quite rare however, since most file names are longer than 3 bytes, and shorter than 15 bytes.

The layout of the filesystem is the following:

```
offset     content

       +---+---+---+---+
 0     | - | r | o | m |  \
       +---+---+---+---+       The ASCII representation of those bytes
 4     | 1 | f | s | - |  /    (i.e. "-rom1fs-")
       +---+---+---+---+
 8     |   full size   |       The number of accessible bytes in this fs.
       +---+---+---+---+
12     |    checksum   |       The checksum of the FIRST 512 BYTES.
       +---+---+---+---+
16     | volume name   |       The zero terminated name of the volume,
       :               :       padded to 16 byte boundary.
       +---+---+---+---+
xx     |     file      |
       :    headers    :
```

Every multi byte value (32 bit words, I'll use the longwords term from now on) must be in big endian order.

The first eight bytes identify the filesystem, even for the casual inspector. After that, in the 3rd longword, it contains the number of bytes accessible from the start of this filesystem. The 4th longword is the checksum of the first 512 bytes (or the number of bytes accessible, whichever is smaller). The applied algorithm is the same as in the AFFS filesystem, namely a simple sum of the longwords (assuming bigendian quantities again). For details, please consult the source. This algorithm was chosen because although it's not quite reliable, it does not require any tables, and it is very simple.

The following bytes are now part of the file system; each file header must begin on a 16 byte boundary:

```
offset     content

       +---+---+---+---+
 0     | next filehdr|X|       The offset of the next file header
       +---+---+---+---+         (zero if no more files)
 4     |   spec.info   |       Info for directories/hard links/devices
       +---+---+---+---+
 8     |     size      |       The size of this file in bytes
       +---+---+---+---+
12     |   checksum    |       Covering the meta data, including the file
       +---+---+---+---+         name, and padding
16     | file name     |       The zero terminated name of the file,
       :               :       padded to 16 byte boundary
       +---+---+---+---+
xx     | file data     |
```

| : | : |
|---|---|

Since the file headers begin always at a 16 byte boundary, the lowest 4 bits would be always zero in the next filehdr pointer. These four bits are used for the mode information. Bits 0..2 specify the type of the file; while bit 4 shows if the file is executable or not. The permissions are assumed to be world readable, if this bit is not set, and world executable if it is; except the character and block devices, they are never accessible for other than owner. The owner of every file is user and group 0, this should never be a problem for the intended use. The mapping of the 8 possible values to file types is the following:

| | | |
|---|---|---|
| 0 | hard link | link destination [file header] |
| 1 | directory | first file's header |
| 2 | regular file | unused, must be zero [MBZ] |
| 3 | symbolic link | unused, MBZ (file data is the link content) |
| 4 | block device | 16/16 bits major/minor number |
| 5 | char device | • ” - |
| 6 | socket | unused, MBZ |
| 7 | fifo | unused, MBZ |

Note that hard links are specifically marked in this filesystem, but they will behave as you can expect (i.e. share the inode number). Note also that it is your responsibility to not create hard link loops, and creating all the . and .. links for directories. This is normally done correctly by the genromfs program. Please refrain from using the executable bits for special purposes on the socket and fifo special files, they may have other uses in the future. Additionally, please remember that only regular files, and symlinks are supposed to have a nonzero size field; they contain the number of bytes available directly after the (padded) file name.

Another thing to note is that romfs works on file headers and data aligned to 16 byte boundaries, but most hardware devices and the block device drivers are unable to cope with smaller than block-sized data. To overcome this limitation, the whole size of the file system must be padded to an 1024 byte boundary.

If you have any problems or suggestions concerning this file system, please contact me. However, think twice before wanting me to add features and code, because the primary and most important advantage of this file system is the small code. On the other hand, don't be alarmed, I'm not getting that much romfs related mail. Now I can understand why Avery wrote poems in the ARCnet docs to get some more feedback. :)

romfs has also a mailing list, and to date, it hasn't received any traffic, so you are welcome to join it to discuss your ideas. :)

It's run by ezmlm, so you can subscribe to it by sending a message to romfs-subscribe@shadow.banki.hu, the content is irrelevant.

Pending issues:

- Permissions and owner information are pretty essential features of a Un*x like system, but romfs does not provide the full possibilities. I have never found this limiting, but others

might.

- The file system is read only, so it can be very small, but in case one would want to write _anything_ to a file system, he still needs a writable file system, thus negating the size advantages. Possible solutions: implement write access as a compile-time option, or a new, similarly small writable filesystem for RAM disks.
- Since the files are only required to have alignment on a 16 byte boundary, it is currently possibly suboptimal to read or execute files from the filesystem. It might be resolved by reordering file data to have most of it (i.e. except the start and the end) laying at "natural" boundaries, thus it would be possible to directly map a big portion of the file contents to the mm subsystem.
- Compression might be an useful feature, but memory is quite a limiting factor in my eyes.
- Where it is used?
- Does it work on other architectures than intel and motorola?

Have fun,

Janos Farkas <chexum@shadow.banki.hu>

## 3.48 CIFS

### 3.48.1 KSMBD - SMB3 Kernel Server

KSMBD is a linux kernel server which implements SMB3 protocol in kernel space for sharing files over network.

#### KSMBD architecture

The subset of performance related operations belong in kernelspace and the other subset which belong to operations which are not really related with performance in userspace. So, DCE/RPC management that has historically resulted into number of buffer overflow issues and dangerous security bugs and user account management are implemented in user space as ksmbd.mountd. File operations that are related with performance (open/read/write/close etc.) in kernel space (ksmbd). This also allows for easier integration with VFS interface for all file operations.

#### ksmbd (kernel daemon)

When the server daemon is started, It starts up a forker thread (ksmbd/interface name) at initialization time and open a dedicated port 445 for listening to SMB requests. Whenever new clients make request, Forker thread will accept the client connection and fork a new thread for dedicated communication channel between the client and the server. It allows for parallel processing of SMB requests(commands) from clients as well as allowing for new clients to make new connections. Each instance is named ksmbd/1~n(port number) to indicate connected clients. Depending on the SMB request types, each new thread can decide to pass through the commands to the user space (ksmbd.mountd), currently DCE/RPC commands are identified to be handled through the user space. To further utilize the linux kernel, it has been chosen to process the commands as workitems and to be executed in the handlers of the ksmbd-io

kworker threads. It allows for multiplexing of the handlers as the kernel take care of initiating extra worker threads if the load is increased and vice versa, if the load is decreased it destroys the extra worker threads. So, after connection is established with client. Dedicated ksmbd/1..n(port number) takes complete ownership of receiving/parsing of SMB commands. Each received command is worked in parallel i.e., There can be multiple clients commands which are worked in parallel. After receiving each command a separated kernel workitem is prepared for each command which is further queued to be handled by ksmbd-io kworkers. So, each SMB workitem is queued to the kworkers. This allows the benefit of load sharing to be managed optimally by the default kernel and optimizing client performance by handling client commands in parallel.

### ksmbd.mountd (user space daemon)

ksmbd.mountd is userspace process to, transfer user account and password that are registered using ksmbd.adduser (part of utils for user space). Further it allows sharing information parameters that parsed from smb.conf to ksmbd in kernel. For the execution part it has a daemon which is continuously running and connected to the kernel interface using netlink socket, it waits for the requests (dcerpc and share/user info). It handles RPC calls (at a minimum few dozen) that are most important for file server from NetShareEnum and NetServerGetInfo. Complete DCE/RPC response is prepared from the user space and passed over to the associated kernel thread for the client.

### KSMBD Feature Status

| Feature name | Status |
|---|---|
| Dialects | Supported. SMB2.1 SMB3.0, SMB3.1.1 dialects (intentionally excludes security vulnerable SMB1 dialect). |
| Auto Negotiation | Supported. |
| Compound Request | Supported. |
| Oplock Cache Mechanism | Supported. |
| SMB2 leases(v1 lease) | Supported. |
| Directory leases(v2 lease) | Supported. |
| Multi-credits | Supported. |
| NTLM/NTLMv2 | Supported. |
| HMAC-SHA256 Signing | Supported. |
| Secure negotiate | Supported. |
| Signing Update | Supported. |
| Pre-authentication integrity | Supported. |
| SMB3 encryption(CCM, GCM) | Supported. (CCM/GCM128 and CCM/GCM256 supported) |
| SMB direct(RDMA) | Supported. |
| SMB3 Multi-channel | Partially Supported. Planned to implement replay/retry mechanisms for future. |
| Receive Side Scaling mode | Supported. |
| SMB3.1.1 POSIX extension | Supported. |
| ACLs | Partially Supported. only DACLs available, SACLs (auditing) is planned for the future. For ownership (SIDs) ksmbd generates random subauth values(then store it to disk) and use uid/gid get from inode as RID for local domain SID. The current acl implementation is limited to standalone server, not a domain member. Integration with Samba tools is being worked on to allow future support for running as a domain member. |
| Kerberos | Supported. |
| Durable handle v1,v2 | Planned for future. |
| Persistent handle | Planned for future. |
| SMB2 notify | Planned for future. |
| Sparse file support | Supported. |
| DCE/RPC support | Partially Supported. a few calls(NetShareEnumAll, NetServerGetInfo, SAMR, LSARPC) that are needed for file server handled via netlink interface from ksmbd.mountd. Additional integration with Samba tools and libraries via upcall is being investigated to allow support for additional DCE/RPC management calls (and future support for Witness protocol e.g.) |
| ksmbd/nfsd interoperability | Planned for future. The features that ksmbd support are Leases, Notify, ACLs and Share modes. |
| SMB3.1.1 Compression | Planned for future. |
| SMB3.1.1 over QUIC | Planned for future. |
| Signing/Encryption over RDMA | Planned for future. |
| SMB3.1.1 GMAC signing support | Planned for future. |

### How to run

1. Download ksmbd-tools(https://github.com/cifsd-team/ksmbd-tools/releases) and compile them.

   - Refer README(https://github.com/cifsd-team/ksmbd-tools/blob/master/README.md) to know how to use ksmbd.mountd/adduser/addshare/control utils

     $ ./autogen.sh $ ./configure --with-rundir=/run $ make && sudo make install

2. Create /usr/local/etc/ksmbd/ksmbd.conf file, add SMB share in ksmbd.conf file.

   - Refer ksmbd.conf.example in ksmbd-utils, See ksmbd.conf manpage for details to configure shares.

     $ man ksmbd.conf

3. Create user/password for SMB share.

   - See ksmbd.adduser manpage.

     $ man ksmbd.adduser $ sudo ksmbd.adduser -a <Enter USERNAME for SMB share access>

4. Insert ksmbd.ko module after build your kernel. No need to load module if ksmbd is built into the kernel.

   - **Set ksmbd in menuconfig(e.g. $ make menuconfig)**

     **[*] Network File Systems --->**

     <M> SMB3 server support (EXPERIMENTAL)

     $ sudo modprobe ksmbd.ko

5. Start ksmbd user space daemon

   $ sudo ksmbd.mountd

6. Access share from Windows or Linux using SMB3 client (cifs.ko or smbclient of samba)

### Shutdown KSMBD

1. **kill user and kernel space daemon**
   # sudo ksmbd.control -s

### How to turn debug print on

Each layer /sys/class/ksmbd-control/debug

1. **Enable all component prints**
   # sudo ksmbd.control -d "all"
2. **Enable one of components (smb, auth, vfs, oplock, ipc, conn, rdma)**
   # sudo ksmbd.control -d "smb"
3. **Show what prints are enabled.**

   **# cat /sys/class/ksmbd-control/debug**
   [smb] auth vfs oplock ipc conn [rdma]

4. **Disable prints:**
   If you try the selected component once more, It is disabled without brackets.

### 3.48.2 Mounting root file system via SMB (cifs.ko)

Written 2019 by Paulo Alcantara <palcantara@suse.de>

Written 2019 by Aurelien Aptel <aaptel@suse.com>

The CONFIG_CIFS_ROOT option enables experimental root file system support over the SMB protocol via cifs.ko.

It introduces a new kernel command-line option called 'cifsroot=' which will tell the kernel to mount the root file system over the network by utilizing SMB or CIFS protocol.

In order to mount, the network stack will also need to be set up by using 'ip=' config option. For more details, see Documentation/admin-guide/nfs/nfsroot.rst.

A CIFS root mount currently requires the use of SMB1+UNIX Extensions which is only supported by the Samba server. SMB1 is the older deprecated version of the protocol but it has been extended to support POSIX features (See [1]). The equivalent extensions for the newer recommended version of the protocol (SMB3) have not been fully implemented yet which means SMB3 doesn't support some required POSIX file system objects (e.g. block devices, pipes, sockets).

As a result, a CIFS root will default to SMB1 for now but the version to use can nonetheless be changed via the 'vers=' mount option. This default will change once the SMB3 POSIX extensions are fully implemented.

#### Server configuration

To enable SMB1+UNIX extensions you will need to set these global settings in Samba smb.conf:

```
[global]
server min protocol = NT1
unix extension = yes        # default
```

#### Kernel command line

```
root=/dev/cifs
```

This is just a virtual device that basically tells the kernel to mount the root file system via SMB protocol.

```
cifsroot=//<server-ip>/<share>[,options]
```

Enables the kernel to mount the root file system via SMB that are located in the <server-ip> and <share> specified in this option.

The default mount options are set in fs/smb/client/cifsroot.c.

**server-ip**
IPv4 address of the server.

**share**
Path to SMB share (rootfs).

**options**
Optional mount options. For more information, see mount.cifs(8).

#### Examples

Export root file system as a Samba share in smb.conf file:

```
...
[linux]
        path = /path/to/rootfs
        read only = no
        guest ok = yes
        force user = root
        force group = root
        browseable = yes
        writeable = yes
        admin users = root
        public = yes
        create mask = 0777
        directory mask = 0777
...
```

Restart smb service:

```
# systemctl restart smb
```

Test it under QEMU on a kernel built with CONFIG_CIFS_ROOT and CONFIG_IP_PNP options enabled:

```
# qemu-system-x86_64 -enable-kvm -cpu host -m 1024 \
-kernel /path/to/linux/arch/x86/boot/bzImage -nographic \
-append "root=/dev/cifs rw ip=dhcp cifsroot=//10.0.2.2/linux,username=foo,
 ↪password=bar console=ttyS0 3"
```

1: https://wiki.samba.org/index.php/UNIX_Extensions

## 3.49 SPU Filesystem

### 3.49.1 spufs

#### Name

spufs - the SPU file system

## Description

The SPU file system is used on PowerPC machines that implement the Cell Broadband Engine Architecture in order to access Synergistic Processor Units (SPUs).

The file system provides a name space similar to posix shared memory or message queues. Users that have write permissions on the file system can use spu_create(2) to establish SPU contexts in the spufs root.

Every SPU context is represented by a directory containing a predefined set of files. These files can be used for manipulating the state of the logical SPU. Users can change permissions on those files, but not actu- ally add or remove files.

## Mount Options

**uid=<uid>**
: set the user owning the mount point, the default is 0 (root).

**gid=<gid>**
: set the group owning the mount point, the default is 0 (root).

## Files

The files in spufs mostly follow the standard behavior for regular sys- tem calls like read(2) or write(2), but often support only a subset of the operations supported on regular file systems. This list details the supported operations and the deviations from the behaviour in the respective man pages.

All files that support the read(2) operation also support readv(2) and all files that support the write(2) operation also support writev(2). All files support the access(2) and stat(2) family of operations, but only the st_mode, st_nlink, st_uid and st_gid fields of struct stat contain reliable information.

All files support the chmod(2)/fchmod(2) and chown(2)/fchown(2) operations, but will not be able to grant permissions that contradict the possible operations, e.g. read access on the wbox file.

The current set of files is:

**/mem**
: the contents of the local storage memory of the SPU. This can be accessed like a regular shared memory file and contains both code and data in the address space of the SPU. The possible operations on an open mem file are:

  **read(2), pread(2), write(2), pwrite(2), lseek(2)**
  : These operate as documented, with the exception that seek(2), write(2) and pwrite(2) are not supported beyond the end of the file. The file size is the size of the local storage of the SPU, which normally is 256 kilobytes.

  **mmap(2)**
  : Mapping mem into the process address space gives access to the SPU local storage within the process ad-

dress space. Only MAP_SHARED mappings are allowed.

**/mbox** The first SPU to CPU communication mailbox. This file is read-only and can be read in units of 32 bits. The file can only be used in non- blocking mode and it even poll() will not block on it. The possible operations on an open mbox file are:

**read(2)**
If a count smaller than four is requested, read returns -1 and sets errno to EINVAL. If there is no data available in the mail box, the return value is set to -1 and errno becomes EAGAIN. When data has been read successfully, four bytes are placed in the data buffer and the value four is returned.

**/ibox** The second SPU to CPU communication mailbox. This file is similar to the first mailbox file, but can be read in blocking I/O mode, and the poll family of system calls can be used to wait for it. The possible operations on an open ibox file are:

**read(2)**
If a count smaller than four is requested, read returns -1 and sets errno to EINVAL. If there is no data available in the mail box and the file descriptor has been opened with O_NONBLOCK, the return value is set to -1 and errno becomes EAGAIN.

If there is no data available in the mail box and the file descriptor has been opened without O_NONBLOCK, the call will block until the SPU writes to its interrupt mailbox channel. When data has been read successfully, four bytes are placed in the data buffer and the value four is returned.

**poll(2)**
Poll on the ibox file returns (POLLIN | POLLRDNORM) whenever data is available for reading.

**/wbox** The CPU to SPU communation mailbox. It is write-only and can be written in units of 32 bits. If the mailbox is full, write() will block and poll can be used to wait for it becoming empty again. The possible operations on an open wbox file are: write(2) If a count smaller than four is requested, write returns -1 and sets errno to EINVAL. If there is no space available in the mail box and the file descriptor has been opened with O_NONBLOCK, the return value is set to -1 and errno becomes EAGAIN.

If there is no space available in the mail box and the file descriptor has been opened without O_NONBLOCK, the call will block until the SPU reads from its PPE mailbox channel. When data has been read success- fully, four bytes are

placed in the data buffer and the value four is returned.

**poll(2)**
Poll on the ibox file returns (POLLOUT | POLLWRNORM) whenever space is available for writing.

**/mbox_stat, /ibox_stat, /wbox_stat** Read-only files that contain the length of the current queue, i.e. how many words can be read from mbox or ibox or how many words can be written to wbox without blocking. The files can be read only in 4-byte units and return a big-endian binary integer number. The possible operations on an open *`box_stat` file are:

**read(2)**
If a count smaller than four is requested, read returns -1 and sets errno to EINVAL. Otherwise, a four byte value is placed in the data buffer, containing the number of elements that can be read from (for mbox_stat and ibox_stat) or written to (for wbox_stat) the respective mail box without blocking or resulting in EAGAIN.

**/npc, /decr, /decr_status, /spu_tag_mask, /event_mask, /srr0** Internal registers of the SPU. The representation is an ASCII string with the numeric value of the next instruction to be executed. These can be used in read/write mode for debugging, but normal operation of programs should not rely on them because access to any of them except npc requires an SPU context save and is therefore very inefficient.

The contents of these files are:

| | |
|---|---|
| npc | Next Program Counter |
| decr | SPU Decrementer |
| decr_status | Decrementer Status |
| spu_tag_mask | MFC tag mask for SPU DMA |
| event_mask | Event mask for SPU interrupts |
| srr0 | Interrupt Return address register |

The possible operations on an open npc, decr, decr_status, spu_tag_mask, event_mask or srr0 file are:

**read(2)**
When the count supplied to the read call is shorter than the required length for the pointer value plus a newline character, subsequent reads from the same file descriptor will result in completing the string, regardless of changes to the register by a running SPU task. When a complete string has been read, all subsequent read operations will return zero bytes and a new file descriptor needs to be opened to read the value again.

**write(2)**
A write operation on the file results in setting the regis-

ter to the value given in the string. The string is parsed from the beginning to the first non-numeric character or the end of the buffer. Subsequent writes to the same file descriptor overwrite the previous setting.

**/fpcr** This file gives access to the Floating Point Status and Control Regis- ter as a four byte long file. The operations on the fpcr file are:

**read(2)**
If a count smaller than four is requested, read returns -1 and sets errno to EINVAL. Otherwise, a four byte value is placed in the data buffer, containing the current value of the fpcr regis- ter.

**write(2)**
If a count smaller than four is requested, write returns -1 and sets errno to EINVAL. Otherwise, a four byte value is copied from the data buffer, updating the value of the fpcr register.

**/signal1, /signal2** The two signal notification channels of an SPU. These are read-write files that operate on a 32 bit word. Writing to one of these files triggers an interrupt on the SPU. The value written to the signal files can be read from the SPU through a channel read or from host user space through the file. After the value has been read by the SPU, it is reset to zero. The possible operations on an open signal1 or sig- nal2 file are:

**read(2)**
If a count smaller than four is requested, read returns -1 and sets errno to EINVAL. Otherwise, a four byte value is placed in the data buffer, containing the current value of the specified signal notification register.

**write(2)**
If a count smaller than four is requested, write returns -1 and sets errno to EINVAL. Otherwise, a four byte value is copied from the data buffer, updating the value of the specified signal notification register. The signal notification register will either be replaced with the input data or will be updated to the bitwise OR of the old value and the input data, depending on the contents of the signal1_type, or signal2_type respectively, file.

**/signal1_type, /signal2_type** These two files change the behavior of the signal1 and signal2 notifi- cation files. The contain a numerical ASCII string which is read as either ”1” or ”0”. In mode 0 (overwrite), the hardware replaces the contents of the signal channel with the data that is written to it. in mode 1 (logical OR), the hardware accumulates the bits that are subse- quently written to it. The possible operations on an open signal1_type or signal2_type file are:

**read(2)**
When the count supplied to the read call is shorter than the required length for the digit plus a newline character, subse- quent reads from the same file descriptor will result in com- pleting the string. When a complete string has been read, all subsequent read operations will return zero bytes and a new file descriptor needs to be opened to read the value again.

**write(2)**
A write operation on the file results in setting the register to the value given in the string. The string is parsed from the beginning to the first non-numeric character or the end of the buffer. Subsequent writes to the same file descriptor overwrite the previous setting.

### Examples

**/etc/fstab entry**
none /spu spufs gid=spu 0 0

### Authors

Arnd Bergmann <arndb@de.ibm.com>, Mark Nutter <mnutter@us.ibm.com>, Ulrich Weigand <Ulrich.Weigand@de.ibm.com>

### See Also

capabilities(7), close(2), spu_create(2), spu_run(2), spufs(7)

## 3.49.2 spu_create

### Name

spu_create - create a new spu context

### Synopsis

```
#include <sys/types.h>
#include <sys/spu.h>

int spu_create(const char *pathname, int flags, mode_t mode);
```

## Description

The spu_create system call is used on PowerPC machines that implement the Cell Broadband Engine Architecture in order to access Synergistic Processor Units (SPUs). It creates a new logical context for an SPU in pathname and returns a handle to associated with it. pathname must point to a non-existing directory in the mount point of the SPU file system (spufs). When spu_create is successful, a directory gets cre- ated on pathname and it is populated with files.

The returned file handle can only be passed to spu_run(2) or closed, other operations are not defined on it. When it is closed, all associ- ated directory entries in spufs are removed. When the last file handle pointing either inside of the context directory or to this file descriptor is closed, the logical SPU context is destroyed.

The parameter flags can be zero or any bitwise or'd combination of the following constants:

**SPU_RAWIO**
Allow mapping of some of the hardware registers of the SPU into user space. This flag requires the CAP_SYS_RAWIO capability, see capabilities(7).

The mode parameter specifies the permissions used for creating the new directory in spufs. mode is modified with the user's umask(2) value and then used for both the directory and the files contained in it. The file permissions mask out some more bits of mode because they typically support only read or write access. See stat(2) for a full list of the possible mode values.

## Return Value

spu_create returns a new file descriptor. It may return -1 to indicate an error condition and set errno to one of the error codes listed below.

## Errors

**EACCES**
The current user does not have write access on the spufs mount point.

EEXIST An SPU context already exists at the given path name.

**EFAULT pathname is not a valid string pointer in the current address**
space.

EINVAL pathname is not a directory in the spufs mount point.

ELOOP Too many symlinks were found while resolving pathname.

EMFILE The process has reached its maximum open file limit.

**ENAMETOOLONG**
pathname was too long.

ENFILE The system has reached the global open file limit.

ENOENT Part of pathname could not be resolved.

ENOMEM The kernel could not allocate all resources required.

**ENOSPC There are not enough SPU resources available to create a new**
context or the user specific limit for the number of SPU con- texts has been reached.

**ENOSYS the functionality is not provided by the current system, because**
either the hardware does not provide SPUs or the spufs module is not loaded.

**ENOTDIR**
A part of pathname is not a directory.

#### Notes

spu_create is meant to be used from libraries that implement a more abstract interface to SPUs, not to be used from regular applications. See http://www.bsc.es/projects/deepcomputing/linuxoncell/ for the rec- ommended libraries.

#### Files

pathname must point to a location beneath the mount point of spufs. By convention, it gets mounted in /spu.

#### Conforming to

This call is Linux specific and only implemented by the ppc64 architec- ture. Programs using this system call are not portable.

#### Bugs

The code does not yet fully implement all features lined out here.

#### Author

Arnd Bergmann <arndb@de.ibm.com>

#### See Also

capabilities(7), close(2), spu_run(2), spufs(7)

### 3.49.3 spu_run

#### Name

spu_run - execute an spu context

### Synopsis

```
#include <sys/spu.h>

int spu_run(int fd, unsigned int *npc, unsigned int *event);
```

### Description

The spu_run system call is used on PowerPC machines that implement the Cell Broadband Engine Architecture in order to access Synergistic Pro- cessor Units (SPUs). It uses the fd that was returned from spu_cre- ate(2) to address a specific SPU context. When the context gets sched- uled to a physical SPU, it starts execution at the instruction pointer passed in npc.

Execution of SPU code happens synchronously, meaning that spu_run does not return while the SPU is still running. If there is a need to exe- cute SPU code in parallel with other code on either the main CPU or other SPUs, you need to create a new thread of execution first, e.g. using the pthread_create(3) call.

When spu_run returns, the current value of the SPU instruction pointer is written back to npc, so you can call spu_run again without updating the pointers.

event can be a NULL pointer or point to an extended status code that gets filled when spu_run returns. It can be one of the following con- stants:

**SPE_EVENT_DMA_ALIGNMENT**
A DMA alignment error

**SPE_EVENT_SPE_DATA_SEGMENT**
A DMA segmentation error

**SPE_EVENT_SPE_DATA_STORAGE**
A DMA storage error

If NULL is passed as the event argument, these errors will result in a signal delivered to the calling process.

### Return Value

spu_run returns the value of the spu_status register or -1 to indicate an error and set errno to one of the error codes listed below. The spu_status register value contains a bit mask of status codes and optionally a 14 bit code returned from the stop-and-signal instruction on the SPU. The bit masks for the status codes are:

**0x02**
SPU was stopped by stop-and-signal.

**0x04**
SPU was stopped by halt.

**0x08**
SPU is waiting for a channel.

**0x10**
SPU is in single-step mode.

**0x20**
SPU has tried to execute an invalid instruction.

**0x40**
SPU has tried to access an invalid channel.

**0x3fff0000**
The bits masked with this value contain the code returned from stop-and-signal.

There are always one or more of the lower eight bits set or an error code is returned from spu_run.

### Errors

**EAGAIN or EWOULDBLOCK**
fd is in non-blocking mode and spu_run would block.

EBADF fd is not a valid file descriptor.

**EFAULT npc is not a valid pointer or status is neither NULL nor a valid**
pointer.

**EINTR A signal occurred while spu_run was in progress. The npc value**
has been updated to the new program counter value if necessary.

EINVAL fd is not a file descriptor returned from spu_create(2).

**ENOMEM Insufficient memory was available to handle a page fault result-**
ing from an MFC direct memory access.

**ENOSYS the functionality is not provided by the current system, because**
either the hardware does not provide SPUs or the spufs module is not loaded.

### Notes

spu_run is meant to be used from libraries that implement a more abstract interface to SPUs, not to be used from regular applications. See http://www.bsc.es/projects/deepcomputing/linuxoncell/ for the rec- ommended libraries.

### Conforming to

This call is Linux specific and only implemented by the ppc64 architec- ture. Programs using this system call are not portable.

### Bugs

The code does not yet fully implement all features lined out here.

### Author

Arnd Bergmann <arndb@de.ibm.com>

### See Also

capabilities(7), close(2), spu_create(2), spufs(7)

## 3.50 Squashfs 4.0 Filesystem

Squashfs is a compressed read-only filesystem for Linux.

It uses zlib, lz4, lzo, or xz compression to compress files, inodes and directories. Inodes in the system are very small and all blocks are packed to minimise data overhead. Block sizes greater than 4K are supported up to a maximum of 1Mbytes (default block size 128K).

Squashfs is intended for general read-only filesystem use, for archival use (i.e. in cases where a .tar.gz file may be used), and in constrained block device/memory systems (e.g. embedded systems) where low overhead is needed.

Mailing list: squashfs-devel@lists.sourceforge.net Web site: www.squashfs.org

### 3.50.1 1. Filesystem Features

Squashfs filesystem features versus Cramfs:

| | | |
|---|---|---|
| Max filesystem size | 2^64 | 256 MiB |
| Max file size | ~ 2 TiB | 16 MiB |
| Max files | unlimited | unlimited |
| Max directories | unlimited | unlimited |
| Max entries per directory | unlimited | unlimited |
| Max block size | 1 MiB | 4 KiB |
| Metadata compression | yes | no |
| Directory indexes | yes | no |
| Sparse file support | yes | no |
| Tail-end packing (fragments) | yes | no |
| Exportable (NFS etc.) | yes | no |
| Hard link support | yes | no |
| ”.” and ”..” in readdir | yes | no |
| Real inode numbers | yes | no |
| 32-bit uids/gids | yes | no |
| File creation time | yes | no |
| Xattr support | yes | no |
| ACL support | no | no |

Squashfs compresses data, inodes and directories. In addition, inode and directory data are highly compacted, and packed on byte boundaries. Each compressed inode is on average 8 bytes in length (the exact length varies on file type, i.e. regular file, directory, symbolic link, and block/char device inodes have different sizes).

### 3.50.2 2. Using Squashfs

As squashfs is a read-only filesystem, the mksquashfs program must be used to create populated squashfs filesystems. This and other squashfs utilities can be obtained from http://www.squashfs.org. Usage instructions can be obtained from this site also.

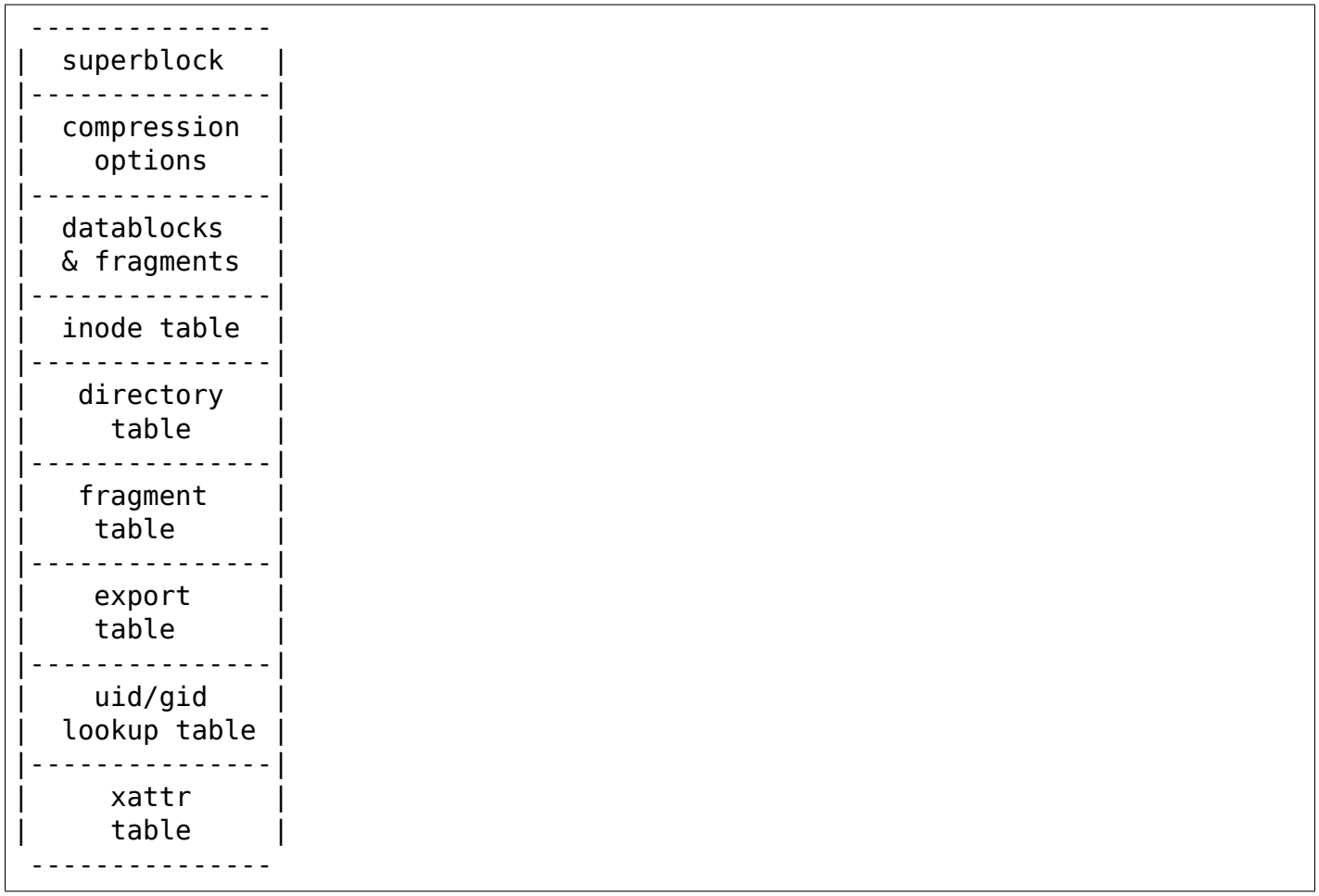
**The squashfs-tools development tree is now located on kernel.org**
git://git.kernel.org/pub/scm/fs/squashfs/squashfs-tools.git

### 3.50.3 2.1 Mount options

<table>
<tr><td>errors=%s</td><td>Specify whether squashfs errors trigger a kernel panic or not<br><table><tr><td>continue</td><td>errors don't trigger a panic (default)</td></tr><tr><td>panic</td><td>trigger a panic when errors are encountered, similar to several other filesystems (e.g. btrfs, ext4, f2fs, GFS2, jfs, ntfs, ubifs) This allows a kernel dump to be saved, useful for analyzing and debugging the corruption.</td></tr></table></td></tr>
<tr><td>threads=%s</td><td>Select the decompression mode or the number of threads<br>If SQUASHFS_CHOICE_DECOMP_BY_MOUNT is set:<br><table><tr><td>single</td><td>use single-threaded decompression (default)<br>Only one block (data or metadata) can be decompressed at any one time. This limits CPU and memory usage to a minimum, but it also gives poor performance on parallel I/O workloads when using multiple CPU machines due to waiting on decompressor availability.</td></tr><tr><td>multi</td><td>use up to two parallel decompressors per core<br>If you have a parallel I/O workload and your system has enough memory, using this option may improve overall I/O performance. It dynamically allocates decompressors on a demand basis.</td></tr><tr><td>percpu</td><td>use a maximum of one decompressor per core<br>It uses percpu variables to ensure decompression is load-balanced across the cores.</td></tr><tr><td>1|2|3|...</td><td>configure the number of threads used for decompression<br>The upper limit is num_online_cpus() * 2.</td></tr></table><br>If SQUASHFS_CHOICE_DECOMP_BY_MOUNT is not set and SQUASHFS_DECOMP_MULTI, SQUASHFS_MOUNT_DECOMP_THREADS are both set:<br><table><tr><td>2|3|...</td><td>configure the number of threads used for decompression</td></tr></table></td></tr>
</table>

### 3.50.4 3. Squashfs Filesystem Design

A squashfs filesystem consists of a maximum of nine parts, packed together on a byte alignment:




Compressed data blocks are written to the filesystem as files are read from the source directory, and checked for duplicates. Once all file data has been written the completed inode, directory, fragment, export, uid/gid lookup and xattr tables are written.

### 3.50.5 3.1 Compression options

Compressors can optionally support compression specific options (e.g. dictionary size). If non-default compression options have been used, then these are stored here.

### 3.50.6 3.2 Inodes

Metadata (inodes and directories) are compressed in 8Kbyte blocks. Each compressed block is prefixed by a two byte length, the top bit is set if the block is uncompressed. A block will be uncompressed if the -noI option is set, or if the compressed block was larger than the uncompressed block.

Inodes are packed into the metadata blocks, and are not aligned to block boundaries, therefore inodes overlap compressed blocks. Inodes are identified by a 48-bit number which encodes the location of the compressed metadata block containing the inode, and the byte offset into that block where the inode is placed (<block, offset>).

To maximise compression there are different inodes for each file type (regular file, directory, device, etc.), the inode contents and length varying with the type.

To further maximise compression, two types of regular file inode and directory inode are defined: inodes optimised for frequently occurring regular files and directories, and extended types where extra information has to be stored.

### 3.50.7 3.3 Directories

Like inodes, directories are packed into compressed metadata blocks, stored in a directory table. Directories are accessed using the start address of the metablock containing the directory and the offset into the decompressed block (<block, offset>).

Directories are organised in a slightly complex way, and are not simply a list of file names. The organisation takes advantage of the fact that (in most cases) the inodes of the files will be in the same compressed metadata block, and therefore, can share the start block. Directories are therefore organised in a two level list, a directory header containing the shared start block value, and a sequence of directory entries, each of which share the shared start block. A new directory header is written once/if the inode start block changes. The directory header/directory entry list is repeated as many times as necessary.

Directories are sorted, and can contain a directory index to speed up file lookup. Directory indexes store one entry per metablock, each entry storing the index/filename mapping to the first directory header in each metadata block. Directories are sorted in alphabetical order, and at lookup the index is scanned linearly looking for the first filename alphabetically larger than the filename being looked up. At this point the location of the metadata block the filename is in has been found. The general idea of the index is to ensure only one metadata block needs to be decompressed to do a lookup irrespective of the length of the directory. This scheme has the advantage that it doesn't require extra memory overhead and doesn't require much extra storage on disk.

### 3.50.8 3.4 File data

Regular files consist of a sequence of contiguous compressed blocks, and/or a compressed fragment block (tail-end packed block). The compressed size of each datablock is stored in a block list contained within the file inode.

To speed up access to datablocks when reading 'large' files (256 Mbytes or larger), the code implements an index cache that caches the mapping from block index to datablock location on disk.

The index cache allows Squashfs to handle large files (up to 1.75 TiB) while retaining a simple and space-efficient block list on disk. The cache is split into slots, caching up to eight 224 GiB files (128 KiB blocks). Larger files use multiple slots, with 1.75 TiB files using all 8 slots. The index cache is designed to be memory efficient, and by default uses 16 KiB.

### 3.50.9 3.5 Fragment lookup table

Regular files can contain a fragment index which is mapped to a fragment location on disk and compressed size using a fragment lookup table. This fragment lookup table is itself stored compressed into metadata blocks. A second index table is used to locate these. This second index table for speed of access (and because it is small) is read at mount time and cached in memory.

### 3.50.10 3.6 Uid/gid lookup table

For space efficiency regular files store uid and gid indexes, which are converted to 32-bit uids/gids using an id look up table. This table is stored compressed into metadata blocks. A second index table is used to locate these. This second index table for speed of access (and because it is small) is read at mount time and cached in memory.

### 3.50.11 3.7 Export table

To enable Squashfs filesystems to be exportable (via NFS etc.) filesystems can optionally (disabled with the -no-exports Mksquashfs option) contain an inode number to inode disk location lookup table. This is required to enable Squashfs to map inode numbers passed in filehandles to the inode location on disk, which is necessary when the export code reinstantiates expired/flushed inodes.

This table is stored compressed into metadata blocks. A second index table is used to locate these. This second index table for speed of access (and because it is small) is read at mount time and cached in memory.

### 3.50.12 3.8 Xattr table

The xattr table contains extended attributes for each inode. The xattrs for each inode are stored in a list, each list entry containing a type, name and value field. The type field encodes the xattr prefix (”user.”, ”trusted.” etc) and it also encodes how the name/value fields should be interpreted. Currently the type indicates whether the value is stored inline (in which case the value field contains the xattr value), or if it is stored out of line (in which case the value field stores a reference to where the actual value is stored). This allows large values to be stored out of line improving scanning and lookup performance and it also allows values to be de-duplicated, the value being stored once, and all other occurrences holding an out of line reference to that value.

The xattr lists are packed into compressed 8K metadata blocks. To reduce overhead in inodes, rather than storing the on-disk location of the xattr list inside each inode, a 32-bit xattr id is stored. This xattr id is mapped into the location of the xattr list using a second xattr id lookup table.

### 3.50.13 4. TODOs and Outstanding Issues

### 3.50.14 4.1 TODO list

Implement ACL support.

### 3.50.15 4.2 Squashfs Internal Cache

Blocks in Squashfs are compressed. To avoid repeatedly decompressing recently accessed data Squashfs uses two small metadata and fragment caches.

The cache is not used for file datablocks, these are decompressed and cached in the page-cache in the normal way. The cache is used to temporarily cache fragment and metadata blocks which have been read as a result of a metadata (i.e. inode or directory) or fragment access. Because metadata and fragments are packed together into blocks (to gain greater compression) the read of a particular piece of metadata or fragment will retrieve other metadata/fragments which have been packed with it, these because of locality-of-reference may be read in the near future. Temporarily caching them ensures they are available for near future access without requiring an additional read and decompress.

In the future this internal cache may be replaced with an implementation which uses the kernel page cache. Because the page cache operates on page sized units this may introduce additional complexity in terms of locking and associated race conditions.

## 3.51 sysfs - _The_ filesystem for exporting kernel objects

Patrick Mochel <mochel@osdl.org>

Mike Murphy <mamurph@cs.clemson.edu>

**Revised**
16 August 2011

**Original**
10 January 2003

### 3.51.1 What it is

sysfs is a RAM-based filesystem initially based on ramfs. It provides a means to export kernel data structures, their attributes, and the linkages between them to userspace.

sysfs is tied inherently to the kobject infrastructure. Please read Documentation/core-api/kobject.rst for more information concerning the kobject interface.

### 3.51.2 Using sysfs

sysfs is always compiled in if CONFIG_SYSFS is defined. You can access it by doing:

```
mount -t sysfs sysfs /sys
```

### 3.51.3 Directory Creation

For every kobject that is registered with the system, a directory is created for it in sysfs. That directory is created as a subdirectory of the kobject's parent, expressing internal object hierarchies to userspace. Top-level directories in sysfs represent the common ancestors of object hierarchies; i.e. the subsystems the objects belong to.

sysfs internally stores a pointer to the kobject that implements a directory in the kernfs_node object associated with the directory. In the past this kobject pointer has been used by sysfs to do reference counting directly on the kobject whenever the file is opened or closed. With the current sysfs implementation the kobject reference count is only modified directly by the function sysfs_schedule_callback().

### 3.51.4 Attributes

Attributes can be exported for kobjects in the form of regular files in the filesystem. sysfs forwards file I/O operations to methods defined for the attributes, providing a means to read and write kernel attributes.

Attributes should be ASCII text files, preferably with only one value per file. It is noted that it may not be efficient to contain only one value per file, so it is socially acceptable to express an array of values of the same type.

Mixing types, expressing multiple lines of data, and doing fancy formatting of data is heavily frowned upon. Doing these things may get you publicly humiliated and your code rewritten without notice.

An attribute definition is simply:

```
struct attribute {
        char                    *name;
        struct module           *owner;
        umode_t                 mode;
};


int sysfs_create_file(struct kobject * kobj, const struct attribute * attr);
void sysfs_remove_file(struct kobject * kobj, const struct attribute * attr);
```

A bare attribute contains no means to read or write the value of the attribute. Subsystems are encouraged to define their own attribute structure and wrapper functions for adding and removing attributes for a specific object type.

For example, the driver model defines struct device_attribute like:

```
struct device_attribute {
        struct attribute    attr;
        ssize_t (*show)(struct device *dev, struct device_attribute *attr,
                        char *buf);
        ssize_t (*store)(struct device *dev, struct device_attribute *attr,
                        const char *buf, size_t count);
};

int device_create_file(struct device *, const struct device_attribute *);
void device_remove_file(struct device *, const struct device_attribute *);
```

It also defines this helper for defining device attributes:

```
#define DEVICE_ATTR(_name, _mode, _show, _store) \
struct device_attribute dev_attr_##_name = __ATTR(_name, _mode, _show, _store)
```

For example, declaring:

```
static DEVICE_ATTR(foo, S_IWUSR | S_IRUGO, show_foo, store_foo);
```

is equivalent to doing:

```
static struct device_attribute dev_attr_foo = {
        .attr = {
                .name = "foo",
                .mode = S_IWUSR | S_IRUGO,
        },
        .show = show_foo,
        .store = store_foo,
};
```

Note as stated in include/linux/kernel.h "OTHER_WRITABLE? Generally considered a bad idea." so trying to set a sysfs file writable for everyone will fail reverting to RO mode for "Others".

For the common cases sysfs.h provides convenience macros to make defining attributes easier as well as making code more concise and readable. The above case could be shortened to:

static struct device_attribute dev_attr_foo = __ATTR_RW(foo);

the list of helpers available to define your wrapper function is:

**__ATTR_RO(name):**
assumes default name_show and mode 0444

**__ATTR_WO(name):**
assumes a name_store only and is restricted to mode 0200 that is root write access only.

**__ATTR_RO_MODE(name, mode):**
for more restrictive RO access; currently only use case is the EFI System Resource Table (see drivers/firmware/efi/esrt.c)

**__ATTR_RW(name):**
assumes default name_show, name_store and setting mode to 0644.

**__ATTR_NULL:**
which sets the name to NULL and is used as end of list indicator (see: kernel/workqueue.c)

### 3.51.5 Subsystem-Specific Callbacks

When a subsystem defines a new attribute type, it must implement a set of sysfs operations for forwarding read and write calls to the show and store methods of the attribute owners:

```
struct sysfs_ops {
        ssize_t (*show)(struct kobject *, struct attribute *, char *);
        ssize_t (*store)(struct kobject *, struct attribute *, const char *,␣
↪size_t);
};
```

[ Subsystems should have already defined a struct kobj_type as a descriptor for this type, which is where the sysfs_ops pointer is stored. See the kobject documentation for more information. ]

When a file is read or written, sysfs calls the appropriate method for the type. The method then translates the generic struct kobject and struct attribute pointers to the appropriate pointer types, and calls the associated methods.

To illustrate:

```
#define to_dev_attr(_attr) container_of(_attr, struct device_attribute, attr)

static ssize_t dev_attr_show(struct kobject *kobj, struct attribute *attr,
                             char *buf)
{
        struct device_attribute *dev_attr = to_dev_attr(attr);
        struct device *dev = kobj_to_dev(kobj);
        ssize_t ret = -EIO;

        if (dev_attr->show)
                ret = dev_attr->show(dev, dev_attr, buf);
        if (ret >= (ssize_t)PAGE_SIZE) {
                printk("dev_attr_show: %pS returned bad count\n",
                                dev_attr->show);
        }
        return ret;
}
```

### 3.51.6 Reading/Writing Attribute Data

To read or write attributes, show() or store() methods must be specified when declaring the attribute. The method types should be as simple as those defined for device attributes:

```
ssize_t (*show)(struct device *dev, struct device_attribute *attr, char *buf);
ssize_t (*store)(struct device *dev, struct device_attribute *attr,
                 const char *buf, size_t count);
```

IOW, they should take only an object, an attribute, and a buffer as parameters.

sysfs allocates a buffer of size (PAGE_SIZE) and passes it to the method. sysfs will call the method exactly once for each read or write. This forces the following behavior on the method

implementations:

- On read(2), the show() method should fill the entire buffer. Recall that an attribute should only be exporting one value, or an array of similar values, so this shouldn't be that expensive.

  This allows userspace to do partial reads and forward seeks arbitrarily over the entire file at will. If userspace seeks back to zero or does a pread(2) with an offset of '0' the show() method will be called again, rearmed, to fill the buffer.

- On write(2), sysfs expects the entire buffer to be passed during the first write. sysfs then passes the entire buffer to the store() method. A terminating null is added after the data on stores. This makes functions like sysfs_streq() safe to use.

  When writing sysfs files, userspace processes should first read the entire file, modify the values it wishes to change, then write the entire buffer back.

  Attribute method implementations should operate on an identical buffer when reading and writing values.

Other notes:

- Writing causes the show() method to be rearmed regardless of current file position.
- The buffer will always be PAGE_SIZE bytes in length. On x86, this is 4096.
- show() methods should return the number of bytes printed into the buffer.
- show() should only use *sysfs_emit()* or *sysfs_emit_at()* when formatting the value to be returned to user space.
- store() should return the number of bytes used from the buffer. If the entire buffer has been used, just return the count argument.
- show() or store() can always return errors. If a bad value comes through, be sure to return an error.
- The object passed to the methods will be pinned in memory via sysfs reference counting its embedded object. However, the physical entity (e.g. device) the object represents may not be present. Be sure to have a way to check this, if necessary.

A very simple (and naive) implementation of a device attribute is:

```
static ssize_t show_name(struct device *dev, struct device_attribute *attr,
                         char *buf)
{
        return sysfs_emit(buf, "%s\n", dev->name);
}

static ssize_t store_name(struct device *dev, struct device_attribute *attr,
                          const char *buf, size_t count)
{
        snprintf(dev->name, sizeof(dev->name), "%.*s",
                 (int)min(count, sizeof(dev->name) - 1), buf);
        return count;
}

static DEVICE_ATTR(name, S_IRUGO, show_name, store_name);
```

(Note that the real implementation doesn't allow userspace to set the name for a device.)

### 3.51.7 Top Level Directory Layout

The sysfs directory arrangement exposes the relationship of kernel data structures.

The top level sysfs directory looks like:

```
block/
bus/
class/
dev/
devices/
firmware/
fs/
hypervisor/
kernel/
module/
net/
power/
```

devices/ contains a filesystem representation of the device tree. It maps directly to the internal kernel device tree, which is a hierarchy of struct device.

bus/ contains flat directory layout of the various bus types in the kernel. Each bus's directory contains two subdirectories:

```
devices/
drivers/
```

devices/ contains symlinks for each device discovered in the system that point to the device's directory under root/.

drivers/ contains a directory for each device driver that is loaded for devices on that particular bus (this assumes that drivers do not span multiple bus types).

fs/ contains a directory for some filesystems. Currently each filesystem wanting to export attributes must create its own hierarchy below fs/ (see ./fuse.rst for an example).

module/ contains parameter values and state information for all loaded system modules, for both builtin and loadable modules.

dev/ contains two directories: char/ and block/. Inside these two directories there are symlinks named <major>:<minor>. These symlinks point to the sysfs directory for the given device. /sys/dev provides a quick way to lookup the sysfs interface for a device from the result of a stat(2) operation.

More information on driver-model specific features can be found in Documentation/driver-api/driver-model/.

TODO: Finish this section.

### 3.51.8 Current Interfaces

The following interface layers currently exist in sysfs.

#### devices (include/linux/device.h)

Structure:

```
struct device_attribute {
        struct attribute    attr;
        ssize_t (*show)(struct device *dev, struct device_attribute *attr,
                        char *buf);
        ssize_t (*store)(struct device *dev, struct device_attribute *attr,
                        const char *buf, size_t count);
};
```

Declaring:

```
DEVICE_ATTR(_name, _mode, _show, _store);
```

Creation/Removal:

```
int device_create_file(struct device *dev, const struct device_attribute *␣
↪attr);
void device_remove_file(struct device *dev, const struct device_attribute *␣
↪attr);
```

#### bus drivers (include/linux/device.h)

Structure:

```
struct bus_attribute {
        struct attribute        attr;
        ssize_t (*show)(const struct bus_type *, char * buf);
        ssize_t (*store)(const struct bus_type *, const char * buf, size_t␣
↪count);
};
```

Declaring:

```
static BUS_ATTR_RW(name);
static BUS_ATTR_RO(name);
static BUS_ATTR_WO(name);
```

Creation/Removal:

```
int bus_create_file(struct bus_type *, struct bus_attribute *);
void bus_remove_file(struct bus_type *, struct bus_attribute *);
```

#### device drivers (include/linux/device.h)

Structure:

```
struct driver_attribute {
        struct attribute        attr;
        ssize_t (*show)(struct device_driver *, char * buf);
        ssize_t (*store)(struct device_driver *, const char * buf,
                         size_t count);
};
```

Declaring:

```
DRIVER_ATTR_RO(_name)
DRIVER_ATTR_RW(_name)
```

Creation/Removal:

```
int driver_create_file(struct device_driver *, const struct driver_attribute␣
↪*);
void driver_remove_file(struct device_driver *, const struct driver_attribute␣
↪*);
```

### 3.51.9 Documentation

The sysfs directory structure and the attributes in each directory define an ABI between the kernel and user space. As for any ABI, it is important that this ABI is stable and properly documented. All new sysfs attributes must be documented in Documentation/ABI. See also Documentation/ABI/README for more information.

## 3.52 SystemV Filesystem

**It implements all of**

- Xenix FS,
- SystemV/386 FS,
- Coherent FS.

To install:

- Answer the 'System V and Coherent filesystem support' question with 'y' when configuring the kernel.
- To mount a disk or a partition, use:

  ```
  mount [-r] -t sysv device mountpoint
  ```

  The file system type names:

```
-t sysv
-t xenix
-t coherent
```

may be used interchangeably, but the last two will eventually disappear.

Bugs in the present implementation:

- Coherent FS:
  - The ”free list interleave” n:m is currently ignored.
  - Only file systems with no filesystem name and no pack name are recognized. (See Coherent ”man mkfs” for a description of these features.)
- SystemV Release 2 FS:

  The superblock is only searched in the blocks 9, 15, 18, which corresponds to the beginning of track 1 on floppy disks. No support for this FS on hard disk yet.

These filesystems are rather similar. Here is a comparison with Minix FS:

- Linux fdisk reports on partitions
  - Minix FS 0x81 Linux/Minix
  - Xenix FS ??
  - SystemV FS ??
  - Coherent FS 0x08 AIX bootable
- Size of a block or zone (data allocation unit on disk)
  - Minix FS 1024
  - Xenix FS 1024 (also 512 ??)
  - SystemV FS 1024 (also 512 and 2048)
  - Coherent FS 512
- General layout: all have one boot block, one super block and separate areas for inodes and for directories/data. On SystemV Release 2 FS (e.g. Microport) the first track is reserved and all the block numbers (including the super block) are offset by one track.
- Byte ordering of ”short” (16 bit entities) on disk:
  - Minix FS little endian 0 1
  - Xenix FS little endian 0 1
  - SystemV FS little endian 0 1
  - Coherent FS little endian 0 1

  Of course, this affects only the file system, not the data of files on it!
- Byte ordering of ”long” (32 bit entities) on disk:
  - Minix FS little endian 0 1 2 3
  - Xenix FS little endian 0 1 2 3
  - SystemV FS little endian 0 1 2 3

- Coherent FS PDP-11 2 3 0 1

Of course, this affects only the file system, not the data of files on it!

- Inode on disk: "short", 0 means non-existent, the root dir ino is:

| | |
|---|---|
| Minix FS | 1 |
| Xenix FS, SystemV FS, Coherent FS | 2 |

- Maximum number of hard links to a file:

| | |
|---|---|
| Minix FS | 250 |
| Xenix FS | ?? |
| SystemV FS | ?? |
| Coherent FS | >=10000 |

- Free inode management:
  - **Minix FS**
    a bitmap
  - **Xenix FS, SystemV FS, Coherent FS**
    There is a cache of a certain number of free inodes in the super-block. When it is exhausted, new free inodes are found using a linear search.
- Free block management:
  - **Minix FS**
    a bitmap
  - **Xenix FS, SystemV FS, Coherent FS**
    Free blocks are organized in a "free list". Maybe a misleading term, since it is not true that every free block contains a pointer to the next free block. Rather, the free blocks are organized in chunks of limited size, and every now and then a free block contains pointers to the free blocks pertaining to the next chunk; the first of these contains pointers and so on. The list terminates with a "block number" 0 on Xenix FS and SystemV FS, with a block zeroed out on Coherent FS.
- Super-block location:

| | |
|---|---|
| Minix FS | block 1 = bytes 1024..2047 |
| Xenix FS | block 1 = bytes 1024..2047 |
| SystemV FS | bytes 512..1023 |
| Coherent FS | block 1 = bytes 512..1023 |

- Super-block layout:
  - Minix FS:

```
unsigned short s_ninodes;
unsigned short s_nzones;
unsigned short s_imap_blocks;
unsigned short s_zmap_blocks;
unsigned short s_firstdatazone;
```

```
unsigned short s_log_zone_size;
unsigned long s_max_size;
unsigned short s_magic;
```

- Xenix FS, SystemV FS, Coherent FS:

```
unsigned short s_firstdatazone;
unsigned long  s_nzones;
unsigned short s_fzone_count;
unsigned long  s_fzones[NICFREE];
unsigned short s_finode_count;
unsigned short s_finodes[NICINOD];
char           s_flock;
char           s_ilock;
char           s_modified;
char           s_rdonly;
unsigned long  s_time;
short          s_dinfo[4]; -- SystemV FS only
unsigned long  s_free_zones;
unsigned short s_free_inodes;
short          s_dinfo[4]; -- Xenix FS only
unsigned short s_interleave_m,s_interleave_n; -- Coherent FS only
char           s_fname[6];
char           s_fpack[6];
```

then they differ considerably:

Xenix FS:

```
char           s_clean;
char           s_fill[371];
long           s_magic;
long           s_type;
```

SystemV FS:

```
long           s_fill[12 or 14];
long           s_state;
long           s_magic;
long           s_type;
```

Coherent FS:

```
unsigned long  s_unique;
```

Note that Coherent FS has no magic.

- Inode layout:
  - Minix FS:

```
unsigned short i_mode;
unsigned short i_uid;
```

```
unsigned long  i_size;
unsigned long  i_time;
unsigned char  i_gid;
unsigned char  i_nlinks;
unsigned short i_zone[7+1+1];
```

- Xenix FS, SystemV FS, Coherent FS:

```
unsigned short i_mode;
unsigned short i_nlink;
unsigned short i_uid;
unsigned short i_gid;
unsigned long  i_size;
unsigned char  i_zone[3*(10+1+1+1)];
unsigned long  i_atime;
unsigned long  i_mtime;
unsigned long  i_ctime;
```

- Regular file data blocks are organized as
  - Minix FS:
    - 7 direct blocks
    - 1 indirect block (pointers to blocks)
    - 1 double-indirect block (pointer to pointers to blocks)
  - Xenix FS, SystemV FS, Coherent FS:
    - 10 direct blocks
    - 1 indirect block (pointers to blocks)
    - 1 double-indirect block (pointer to pointers to blocks)
    - 1 triple-indirect block (pointer to pointers to pointers to blocks)

| | | |
|---|---|---|
| Minix FS | 32 | 32 |
| Xenix FS | 64 | 16 |
| SystemV FS | 64 | 16 |
| Coherent FS | 64 | 8 |

- Directory entry on disk
  - Minix FS:

```
unsigned short inode;
char name[14/30];
```

  - Xenix FS, SystemV FS, Coherent FS:

```
unsigned short inode;
char name[14];
```

| | | |
|---|---|---|
| Minix FS | 16/32 | 64/32 |
| Xenix FS | 16 | 64 |
| SystemV FS | 16 | 64 |
| Coherent FS | 16 | 32 |

- How to implement symbolic links such that the host fsck doesn't scream:
  - Minix FS normal
  - Xenix FS kludge: as regular files with chmod 1000
  - SystemV FS ??
  - Coherent FS kludge: as regular files with chmod 1000

Notation: We often speak of a ”block” but mean a zone (the allocation unit) and not the disk driver's notion of ”block”.

## 3.53 Tmpfs

Tmpfs is a file system which keeps all of its files in virtual memory.

Everything in tmpfs is temporary in the sense that no files will be created on your hard drive. If you unmount a tmpfs instance, everything stored therein is lost.

tmpfs puts everything into the kernel internal caches and grows and shrinks to accommodate the files it contains and is able to swap unneeded pages out to swap space, if swap was enabled for the tmpfs mount. tmpfs also supports THP.

tmpfs extends ramfs with a few userspace configurable options listed and explained further below, some of which can be reconfigured dynamically on the fly using a remount ('mount -o remount ...') of the filesystem. A tmpfs filesystem can be resized but it cannot be resized to a size below its current usage. tmpfs also supports POSIX ACLs, and extended attributes for the trusted.*, security.* and user.* namespaces. ramfs does not use swap and you cannot modify any parameter for a ramfs filesystem. The size limit of a ramfs filesystem is how much memory you have available, and so care must be taken if used so to not run out of memory.

An alternative to tmpfs and ramfs is to use brd to create RAM disks (/dev/ram*), which allows you to simulate a block device disk in physical RAM. To write data you would just then need to create an regular filesystem on top this ramdisk. As with ramfs, brd ramdisks cannot swap. brd ramdisks are also configured in size at initialization and you cannot dynamically resize them. Contrary to brd ramdisks, tmpfs has its own filesystem, it does not rely on the block layer at all.

Since tmpfs lives completely in the page cache and optionally on swap, all tmpfs pages will be shown as ”Shmem” in /proc/meminfo and ”Shared” in free(1). Notice that these counters also include shared memory (shmem, see ipcs(1)). The most reliable way to get the count is using df(1) and du(1).

tmpfs has the following uses:

1) There is always a kernel internal mount which you will not see at all. This is used for shared anonymous mappings and SYSV shared memory.

This mount does not depend on CONFIG_TMPFS. If CONFIG_TMPFS is not set, the user visible part of tmpfs is not built. But the internal mechanisms are always present.

2) glibc 2.2 and above expects tmpfs to be mounted at /dev/shm for POSIX shared memory (shm_open, shm_unlink). Adding the following line to /etc/fstab should take care of this:

```
tmpfs   /dev/shm        tmpfs   defaults        0 0
```

Remember to create the directory that you intend to mount tmpfs on if necessary.

This mount is _not_ needed for SYSV shared memory. The internal mount is used for that. (In the 2.3 kernel versions it was necessary to mount the predecessor of tmpfs (shm fs) to use SYSV shared memory.)

3) Some people (including me) find it very convenient to mount it e.g. on /tmp and /var/tmp and have a big swap partition. And now loop mounts of tmpfs files do work, so mkinitrd shipped by most distributions should succeed with a tmpfs /tmp.

4) And probably a lot more I do not know about :-)

tmpfs has three mount options for sizing:

| | |
|---|---|
| size | The limit of allocated bytes for this tmpfs instance. The default is half of your physical RAM without swap. If you oversize your tmpfs instances the machine will deadlock since the OOM handler will not be able to free that memory. |
| nr_blocks | The same as size, but in blocks of PAGE_SIZE. |
| nr_inodes | The maximum number of inodes for this instance. The default is half of the number of your physical RAM pages, or (on a machine with highmem) the number of lowmem RAM pages, whichever is the lower. |

These parameters accept a suffix k, m or g for kilo, mega and giga and can be changed on remount. The size parameter also accepts a suffix % to limit this tmpfs instance to that percentage of your physical RAM: the default, when neither size nor nr_blocks is specified, is size=50%

If nr_blocks=0 (or size=0), blocks will not be limited in that instance; if nr_inodes=0, inodes will not be limited. It is generally unwise to mount with such options, since it allows any user with write access to use up all the memory on the machine; but enhances the scalability of that instance in a system with many CPUs making intensive use of it.

If nr_inodes is not 0, that limited space for inodes is also used up by extended attributes: ”df -i”'s IUsed and IUse% increase, IFree decreases.

tmpfs blocks may be swapped out, when there is a shortage of memory. tmpfs has a mount option to disable its use of swap:

| | |
|---|---|
| noswap | Disables swap. Remounts must respect the original settings. By default swap is enabled. |

tmpfs also supports Transparent Huge Pages which requires a kernel configured with CONFIG_TRANSPARENT_HUGEPAGE and with huge supported for your system (has_transparent_hugepage(), which is architecture specific). The mount options for this are:

| | |
|---|---|
| huge=never | Do not allocate huge pages. This is the default. |
| huge=always | Attempt to allocate huge page every time a new page is needed. |
| huge=within_size | Only allocate huge page if it will be fully within i_size. Also respect madvise(2) hints. |
| huge=advise | Only allocate huge page if requested with madvise(2). |

See also Documentation/admin-guide/mm/transhuge.rst, which describes the sysfs file /sys/kernel/mm/transparent_hugepage/shmem_enabled: which can be used to deny huge pages on all tmpfs mounts in an emergency, or to force huge pages on all tmpfs mounts for testing.

tmpfs also supports quota with the following mount options

| | |
|---|---|
| quota | User and group quota accounting and enforcement is enabled on the mount. Tmpfs is using hidden system quota files that are initialized on mount. |
| usrquota | User quota accounting and enforcement is enabled on the mount. |
| grpquota | Group quota accounting and enforcement is enabled on the mount. |
| usrquota_block_hardlimit | Set global user quota block hard limit. |
| usrquota_inode_hardlimit | Set global user quota inode hard limit. |
| grpquota_block_hardlimit | Set global group quota block hard limit. |
| grpquota_inode_hardlimit | Set global group quota inode hard limit. |

None of the quota related mount options can be set or changed on remount.

Quota limit parameters accept a suffix k, m or g for kilo, mega and giga and can't be changed on remount. Default global quota limits are taking effect for any and all user/group/project except root the first time the quota entry for user/group/project id is being accessed - typically the first time an inode with a particular id ownership is being created after the mount. In other words, instead of the limits being initialized to zero, they are initialized with the particular value provided with these mount options. The limits can be changed for any user/group id at any time as they normally can be.

Note that tmpfs quotas do not support user namespaces so no uid/gid translation is done if quotas are enabled inside user namespaces.

tmpfs has a mount option to set the NUMA memory allocation policy for all files in that instance (if CONFIG_NUMA is enabled) - which can be adjusted on the fly via 'mount -o remount ...'

| | |
|---|---|
| mpol=default | use the process allocation policy (see set_mempolicy(2)) |
| mpol=prefer:Node | prefers to allocate memory from the given Node |
| mpol=bind:NodeList | allocates memory only from nodes in NodeList |
| mpol=interleave | prefers to allocate from each node in turn |
| mpol=interleave:NodeList | allocates from each node of NodeList in turn |
| mpol=local | prefers to allocate memory from the local node |

NodeList format is a comma-separated list of decimal numbers and ranges, a range being two hyphen-separated decimal numbers, the smallest and largest node numbers in the range. For example, mpol=bind:0-3,5,7,9-15

A memory policy with a valid NodeList will be saved, as specified, for use at file creation time. When a task allocates a file in the file system, the mount option memory policy will

be applied with a NodeList, if any, modified by the calling task's cpuset constraints [See Documentation/admin-guide/cgroup-v1/cpusets.rst] and any optional flags, listed below. If the resulting NodeLists is the empty set, the effective memory policy for the file will revert to ”default” policy.

NUMA memory allocation policies have optional flags that can be used in conjunction with their modes. These optional flags can be specified when tmpfs is mounted by appending them to the mode before the NodeList. See Documentation/admin-guide/mm/numa_memory_policy.rst for a list of all available memory allocation policy mode flags and their effect on memory policy.

```
=static         is equivalent to        MPOL_F_STATIC_NODES
=relative       is equivalent to        MPOL_F_RELATIVE_NODES
```

For example, mpol=bind=static:NodeList, is the equivalent of an allocation policy of MPOL_BIND | MPOL_F_STATIC_NODES.

Note that trying to mount a tmpfs with an mpol option will fail if the running kernel does not support NUMA; and will fail if its nodelist specifies a node which is not online. If your system relies on that tmpfs being mounted, but from time to time runs a kernel built without NUMA capability (perhaps a safe recovery kernel), or with fewer nodes online, then it is advisable to omit the mpol option from automatic mount options. It can be added later, when the tmpfs is already mounted on MountPoint, by 'mount -o remount,mpol=Policy:NodeList MountPoint'.

To specify the initial root directory you can use the following mount options:

| | |
|---|---|
| mode | The permissions as an octal number |
| uid | The user id |
| gid | The group id |

These options do not have any effect on remount. You can change these parameters with chmod(1), chown(1) and chgrp(1) on a mounted filesystem.

tmpfs has a mount option to select whether it will wrap at 32- or 64-bit inode numbers:

| | |
|---|---|
| inode64 | Use 64-bit inode numbers |
| inode32 | Use 32-bit inode numbers |

On a 32-bit kernel, inode32 is implicit, and inode64 is refused at mount time. On a 64-bit kernel, CONFIG_TMPFS_INODE64 sets the default. inode64 avoids the possibility of multiple files with the same inode number on a single device; but risks glibc failing with EOVERFLOW once 33-bit inode numbers are reached - if a long-lived tmpfs is accessed by 32-bit applications so ancient that opening a file larger than 2GiB fails with EINVAL.

So 'mount -t tmpfs -o size=10G,nr_inodes=10k,mode=700 tmpfs /mytmpfs' will give you tmpfs instance on /mytmpfs which can allocate 10GB RAM/SWAP in 10240 inodes and it is only accessible by root.

**Author**
Christoph Rohland <cr@sap.com>, 1.12.01

**Updated**
Hugh Dickins, 4 June 2007

**Updated**
KOSAKI Motohiro, 16 Mar 2010

**Updated**
Chris Down, 13 July 2020

## 3.54 UBI File System

### 3.54.1 Introduction

UBIFS file-system stands for UBI File System. UBI stands for ”Unsorted Block Images”. UBIFS is a flash file system, which means it is designed to work with flash devices. It is important to understand, that UBIFS is completely different to any traditional file-system in Linux, like Ext2, XFS, JFS, etc. UBIFS represents a separate class of file-systems which work with MTD devices, not block devices. The other Linux file-system of this class is JFFS2.

To make it more clear, here is a small comparison of MTD devices and block devices.

**1 MTD devices represent flash devices and they consist of eraseblocks of**
rather large size, typically about 128KiB. Block devices consist of small blocks, typically 512 bytes.

**2 MTD devices support 3 main operations - read from some offset within an**
eraseblock, write to some offset within an eraseblock, and erase a whole eraseblock. Block devices support 2 main operations - read a whole block and write a whole block.

**3 The whole eraseblock has to be erased before it becomes possible to**
re-write its contents. Blocks may be just re-written.

**4 Eraseblocks become worn out after some number of erase cycles -**
typically 100K-1G for SLC NAND and NOR flashes, and 1K-10K for MLC NAND flashes. Blocks do not have the wear-out property.

**5 Eraseblocks may become bad (only on NAND flashes) and software should**
deal with this. Blocks on hard drives typically do not become bad, because hardware has mechanisms to substitute bad blocks, at least in modern LBA disks.

It should be quite obvious why UBIFS is very different to traditional file-systems.

UBIFS works on top of UBI. UBI is a separate software layer which may be found in drivers/mtd/ubi. UBI is basically a volume management and wear-leveling layer. It provides so called UBI volumes which is a higher level abstraction than a MTD device. The programming model of UBI devices is very similar to MTD devices - they still consist of large eraseblocks, they have read/write/erase operations, but UBI devices are devoid of limitations like wear and bad blocks (items 4 and 5 in the above list).

In a sense, UBIFS is a next generation of JFFS2 file-system, but it is very different and incompatible to JFFS2. The following are the main differences.

- JFFS2 works on top of MTD devices, UBIFS depends on UBI and works on top of UBI volumes.
- JFFS2 does not have on-media index and has to build it while mounting, which requires full media scan. UBIFS maintains the FS indexing information on the flash media and does not require full media scan, so it mounts many times faster than JFFS2.
- JFFS2 is a write-through file-system, while UBIFS supports write-back, which makes UBIFS much faster on writes.

Similarly to JFFS2, UBIFS supports on-the-fly compression which makes it possible to fit quite a lot of data to the flash.

Similarly to JFFS2, UBIFS is tolerant of unclean reboots and power-cuts. It does not need stuff like fsck.ext2. UBIFS automatically replays its journal and recovers from crashes, ensuring that the on-flash data structures are consistent.

UBIFS scales logarithmically (most of the data structures it uses are trees), so the mount time and memory consumption do not linearly depend on the flash size, like in case of JFFS2. This is because UBIFS maintains the FS index on the flash media. However, UBIFS depends on UBI, which scales linearly. So overall UBI/UBIFS stack scales linearly. Nevertheless, UBI/UBIFS scales considerably better than JFFS2.

The authors of UBIFS believe, that it is possible to develop UBI2 which would scale logarithmically as well. UBI2 would support the same API as UBI, but it would be binary incompatible to UBI. So UBIFS would not need to be changed to use UBI2

### 3.54.2 Mount options

(*) == default.

| | |
|---|---|
| bulk_read | read more in one go to take advantage of flash media that read faster sequentially |
| no_bulk_read (*) | do not bulk-read |
| no_chk_data_crc (*) | skip checking of CRCs on data nodes in order to improve read performance. Use this option only if the flash media is highly reliable. The effect of this option is that corruption of the contents of a file can go unnoticed. |
| chk_data_crc | do not skip checking CRCs on data nodes |
| compr=none | override default compressor and set it to ”none” |
| compr=lzo | override default compressor and set it to ”lzo” |
| compr=zlib | override default compressor and set it to ”zlib” |
| auth_key= | specify the key used for authenticating the filesystem. Passing this option makes authentication mandatory. The passed key must be present in the kernel keyring and must be of type 'logon' |
| auth_hash_name= | The hash algorithm used for authentication. Used for both hashing and for creating HMACs. Typical values include ”sha256” or ”sha512” |

### 3.54.3 Quick usage instructions

The UBI volume to mount is specified using ”ubiX_Y” or ”ubiX:NAME” syntax, where ”X” is UBI device number, ”Y” is UBI volume number, and ”NAME” is UBI volume name.

Mount volume 0 on UBI device 0 to /mnt/ubifs:

```
$ mount -t ubifs ubi0_0 /mnt/ubifs
```

Mount ”rootfs” volume of UBI device 0 to /mnt/ubifs (”rootfs” is volume name):

```
$ mount -t ubifs ubi0:rootfs /mnt/ubifs
```

The following is an example of the kernel boot arguments to attach mtd0 to UBI and mount volume ”rootfs”: ubi.mtd=0 root=ubi0:rootfs rootfstype=ubifs

### 3.54.4 References

UBIFS documentation and FAQ/HOWTO at the MTD web site:

- http://www.linux-mtd.infradead.org/doc/ubifs.html
- http://www.linux-mtd.infradead.org/faq/ubifs.html

## 3.55 UBIFS Authentication Support

### 3.55.1 Introduction

UBIFS utilizes the fscrypt framework to provide confidentiality for file contents and file names. This prevents attacks where an attacker is able to read contents of the filesystem on a single point in time. A classic example is a lost smartphone where the attacker is unable to read personal data stored on the device without the filesystem decryption key.

At the current state, UBIFS encryption however does not prevent attacks where the attacker is able to modify the filesystem contents and the user uses the device afterwards. In such a scenario an attacker can modify filesystem contents arbitrarily without the user noticing. One example is to modify a binary to perform a malicious action when executed [DMC-CBC-ATTACK]. Since most of the filesystem metadata of UBIFS is stored in plain, this makes it fairly easy to swap files and replace their contents.

Other full disk encryption systems like dm-crypt cover all filesystem metadata, which makes such kinds of attacks more complicated, but not impossible. Especially, if the attacker is given access to the device multiple points in time. For dm-crypt and other filesystems that build upon the Linux block IO layer, the dm-integrity or dm-verity subsystems [DM-INTEGRITY, DM-VERITY] can be used to get full data authentication at the block layer. These can also be combined with dm-crypt [CRYPTSETUP2].

This document describes an approach to get file contents _and_ full metadata authentication for UBIFS. Since UBIFS uses fscrypt for file contents and file name encryption, the authentication system could be tied into fscrypt such that existing features like key derivation can be utilized. It should however also be possible to use UBIFS authentication without using encryption.

#### MTD, UBI & UBIFS

On Linux, the MTD (Memory Technology Devices) subsystem provides a uniform interface to access raw flash devices. One of the more prominent subsystems that work on top of MTD is UBI (Unsorted Block Images). It provides volume management for flash devices and is thus somewhat similar to LVM for block devices. In addition, it deals with flash-specific wear-leveling and transparent I/O error handling. UBI offers logical erase blocks (LEBs) to the layers on top of it and maps them transparently to physical erase blocks (PEBs) on the flash.

UBIFS is a filesystem for raw flash which operates on top of UBI. Thus, wear leveling and some flash specifics are left to UBI, while UBIFS focuses on scalability, performance and recoverability.

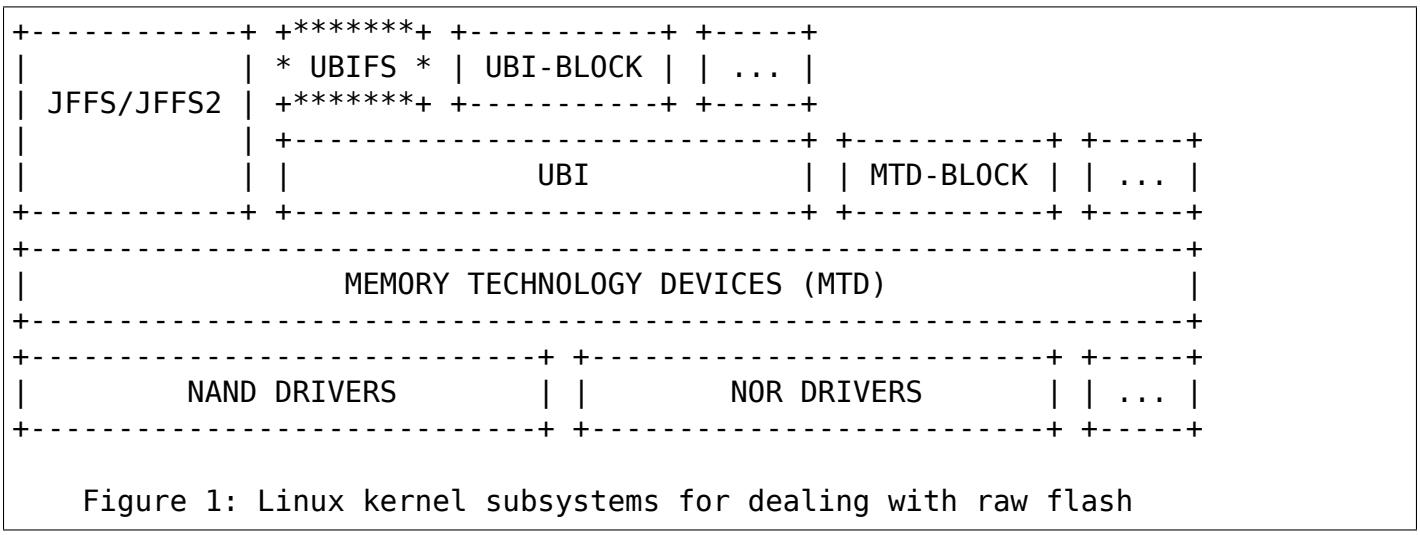



Figure 1: Linux kernel subsystems for dealing with raw flash

Internally, UBIFS maintains multiple data structures which are persisted on the flash:

- *Index*: an on-flash B+ tree where the leaf nodes contain filesystem data
- *Journal*: an additional data structure to collect FS changes before updating the on-flash index and reduce flash wear.
- *Tree Node Cache (TNC)*: an in-memory B+ tree that reflects the current FS state to avoid frequent flash reads. It is basically the in-memory representation of the index, but contains additional attributes.
- *LEB property tree (LPT)*: an on-flash B+ tree for free space accounting per UBI LEB.

In the remainder of this section we will cover the on-flash UBIFS data structures in more detail. The TNC is of less importance here since it is never persisted onto the flash directly. More details on UBIFS can also be found in [UBIFS-WP].

### UBIFS Index & Tree Node Cache

Basic on-flash UBIFS entities are called *nodes*. UBIFS knows different types of nodes. Eg. data nodes (`struct ubifs_data_node`) which store chunks of file contents or inode nodes (`struct ubifs_ino_node`) which represent VFS inodes. Almost all types of nodes share a common header (`ubifs_ch`) containing basic information like node type, node length, a sequence number, etc. (see `fs/ubifs/ubifs-media.h` in kernel source). Exceptions are entries of the LPT and some less important node types like padding nodes which are used to pad unusable content at the end of LEBs.

To avoid re-writing the whole B+ tree on every single change, it is implemented as *wandering tree*, where only the changed nodes are re-written and previous versions of them are obsoleted without erasing them right away. As a result, the index is not stored in a single place on the flash, but *wanders* around and there are obsolete parts on the flash as long as the LEB containing them is not reused by UBIFS. To find the most recent version of the index, UBIFS stores a special node called *master node* into UBI LEB 1 which always points to the most recent root node of the UBIFS index. For recoverability, the master node is additionally duplicated to LEB 2. Mounting UBIFS is thus a simple read of LEB 1 and 2 to get the current master node and from there get the location of the most recent on-flash index.

The TNC is the in-memory representation of the on-flash index. It contains some additional

runtime attributes per node which are not persisted. One of these is a dirty-flag which marks nodes that have to be persisted the next time the index is written onto the flash. The TNC acts as a write-back cache and all modifications of the on-flash index are done through the TNC. Like other caches, the TNC does not have to mirror the full index into memory, but reads parts of it from flash whenever needed. A *commit* is the UBIFS operation of updating the on-flash filesystem structures like the index. On every commit, the TNC nodes marked as dirty are written to the flash to update the persisted index.

### Journal

To avoid wearing out the flash, the index is only persisted (*committed*) when certain conditions are met (eg. `fsync(2)`). The journal is used to record any changes (in form of inode nodes, data nodes etc.) between commits of the index. During mount, the journal is read from the flash and replayed onto the TNC (which will be created on-demand from the on-flash index).

UBIFS reserves a bunch of LEBs just for the journal called *log area*. The amount of log area LEBs is configured on filesystem creation (using `mkfs.ubifs`) and stored in the superblock node. The log area contains only two types of nodes: *reference nodes* and *commit start nodes*. A commit start node is written whenever an index commit is performed. Reference nodes are written on every journal update. Each reference node points to the position of other nodes (inode nodes, data nodes etc.) on the flash that are part of this journal entry. These nodes are called *buds* and describe the actual filesystem changes including their data.

The log area is maintained as a ring. Whenever the journal is almost full, a commit is initiated. This also writes a commit start node so that during mount, UBIFS will seek for the most recent commit start node and just replay every reference node after that. Every reference node before the commit start node will be ignored as they are already part of the on-flash index.

When writing a journal entry, UBIFS first ensures that enough space is available to write the reference node and buds part of this entry. Then, the reference node is written and afterwards the buds describing the file changes. On replay, UBIFS will record every reference node and inspect the location of the referenced LEBs to discover the buds. If these are corrupt or missing, UBIFS will attempt to recover them by re-reading the LEB. This is however only done for the last referenced LEB of the journal. Only this can become corrupt because of a power cut. If the recovery fails, UBIFS will not mount. An error for every other LEB will directly cause UBIFS to fail the mount operation.

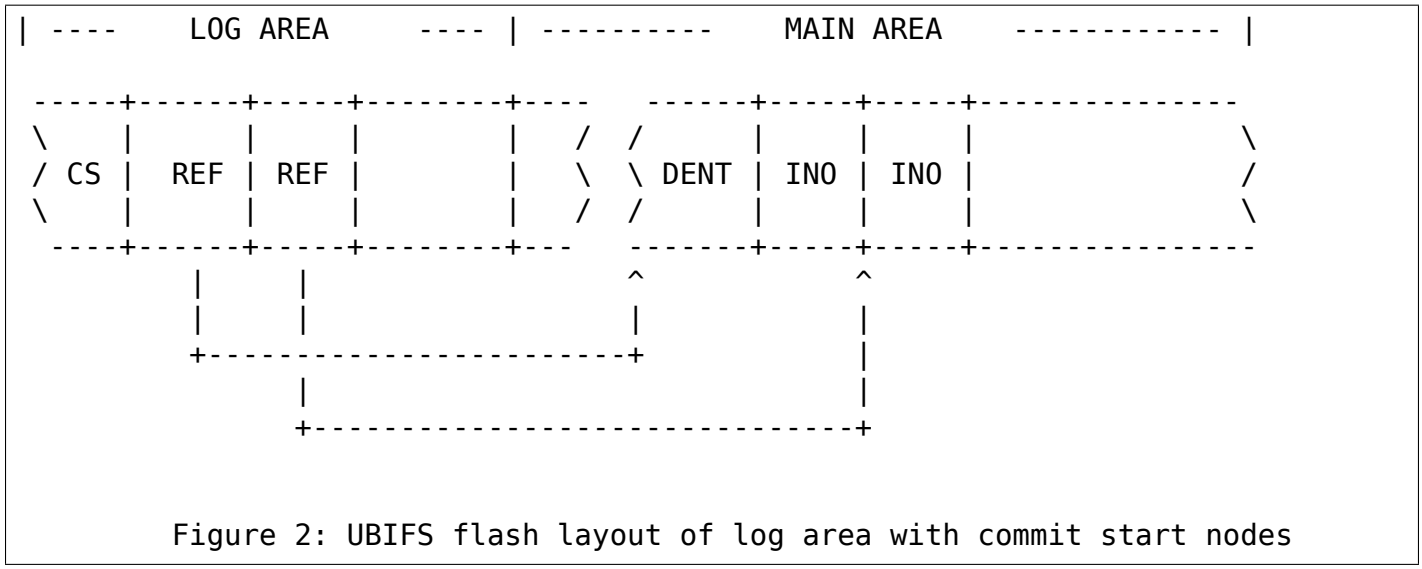



Figure 2: UBIFS flash layout of log area with commit start nodes

```
                (CS) and reference nodes (REF) pointing to main area
                containing their buds
```

#### LEB Property Tree/Table

The LEB property tree is used to store per-LEB information. This includes the LEB type and amount of free and *dirty* (old, obsolete content) space[1] on the LEB. The type is important, because UBIFS never mixes index nodes with data nodes on a single LEB and thus each LEB has a specific purpose. This again is useful for free space calculations. See [UBIFS-WP] for more details.

The LEB property tree again is a B+ tree, but it is much smaller than the index. Due to its smaller size it is always written as one chunk on every commit. Thus, saving the LPT is an atomic operation.

### 3.55.2 UBIFS Authentication

This chapter introduces UBIFS authentication which enables UBIFS to verify the authenticity and integrity of metadata and file contents stored on flash.

#### Threat Model

UBIFS authentication enables detection of offline data modification. While it does not prevent it, it enables (trusted) code to check the integrity and authenticity of on-flash file contents and filesystem metadata. This covers attacks where file contents are swapped.

UBIFS authentication will not protect against rollback of full flash contents. Ie. an attacker can still dump the flash and restore it at a later time without detection. It will also not protect against partial rollback of individual index commits. That means that an attacker is able to partially undo changes. This is possible because UBIFS does not immediately overwrites obsolete versions of the index tree or the journal, but instead marks them as obsolete and garbage collection erases them at a later time. An attacker can use this by erasing parts of the current tree and restoring old versions that are still on the flash and have not yet been erased. This is possible, because every commit will always write a new version of the index root node and the master node without overwriting the previous version. This is further helped by the wear-leveling operations of UBI which copies contents from one physical eraseblock to another and does not atomically erase the first eraseblock.

UBIFS authentication does not cover attacks where an attacker is able to execute code on the device after the authentication key was provided. Additional measures like secure boot and trusted boot have to be taken to ensure that only trusted code is executed on a device.

[1] Since LEBs can only be appended and never overwritten, there is a difference between free space ie. the remaining space left on the LEB to be written to without erasing it and previously written content that is obsolete but can't be overwritten without erasing the full LEB.

### Authentication

To be able to fully trust data read from flash, all UBIFS data structures stored on flash are authenticated. That is:

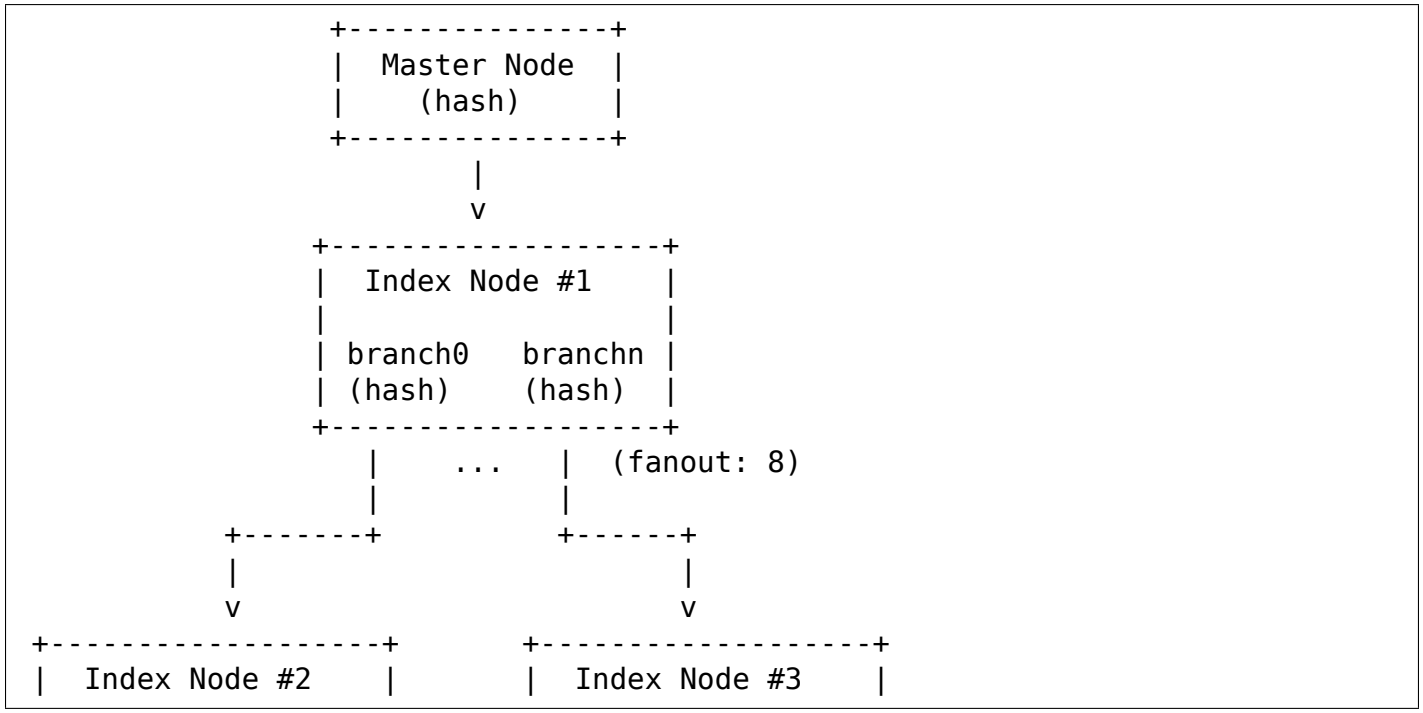
- The index which includes file contents, file metadata like extended attributes, file length etc.
- The journal which also contains file contents and metadata by recording changes to the filesystem
- The LPT which stores UBI LEB metadata which UBIFS uses for free space accounting

### Index Authentication

Through UBIFS' concept of a wandering tree, it already takes care of only updating and persisting changed parts from leaf node up to the root node of the full B+ tree. This enables us to augment the index nodes of the tree with a hash over each node's child nodes. As a result, the index basically also a Merkle tree. Since the leaf nodes of the index contain the actual filesystem data, the hashes of their parent index nodes thus cover all the file contents and file metadata. When a file changes, the UBIFS index is updated accordingly from the leaf nodes up to the root node including the master node. This process can be hooked to recompute the hash only for each changed node at the same time. Whenever a file is read, UBIFS can verify the hashes from each leaf node up to the root node to ensure the node's integrity.

To ensure the authenticity of the whole index, the UBIFS master node stores a keyed hash (HMAC) over its own contents and a hash of the root node of the index tree. As mentioned above, the master node is always written to the flash whenever the index is persisted (ie. on index commit).

Using this approach only UBIFS index nodes and the master node are changed to include a hash. All other types of nodes will remain unchanged. This reduces the storage overhead which is precious for users of UBIFS (ie. embedded devices).



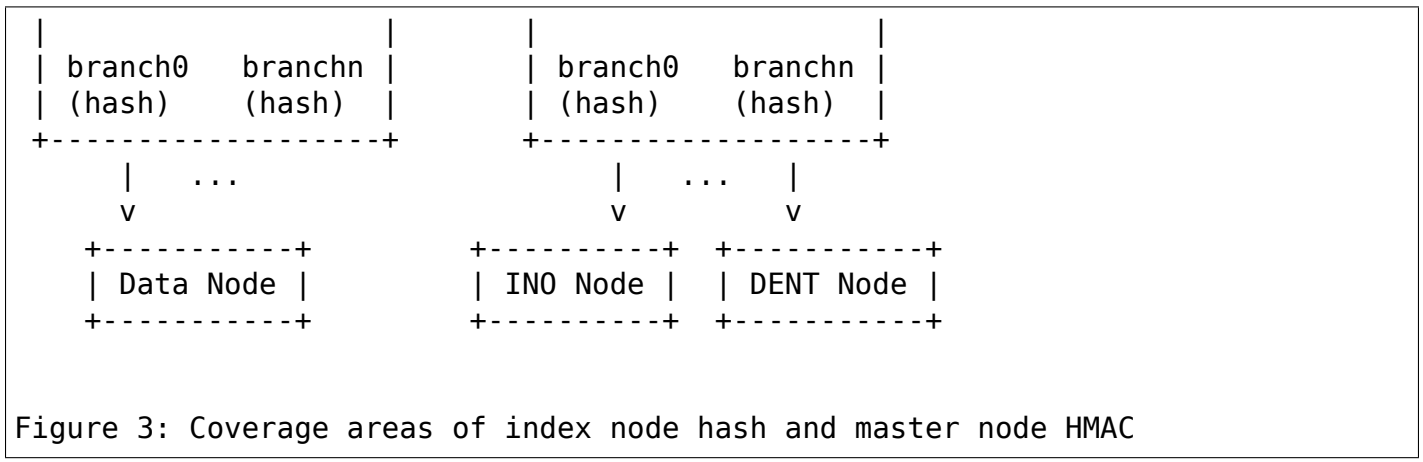


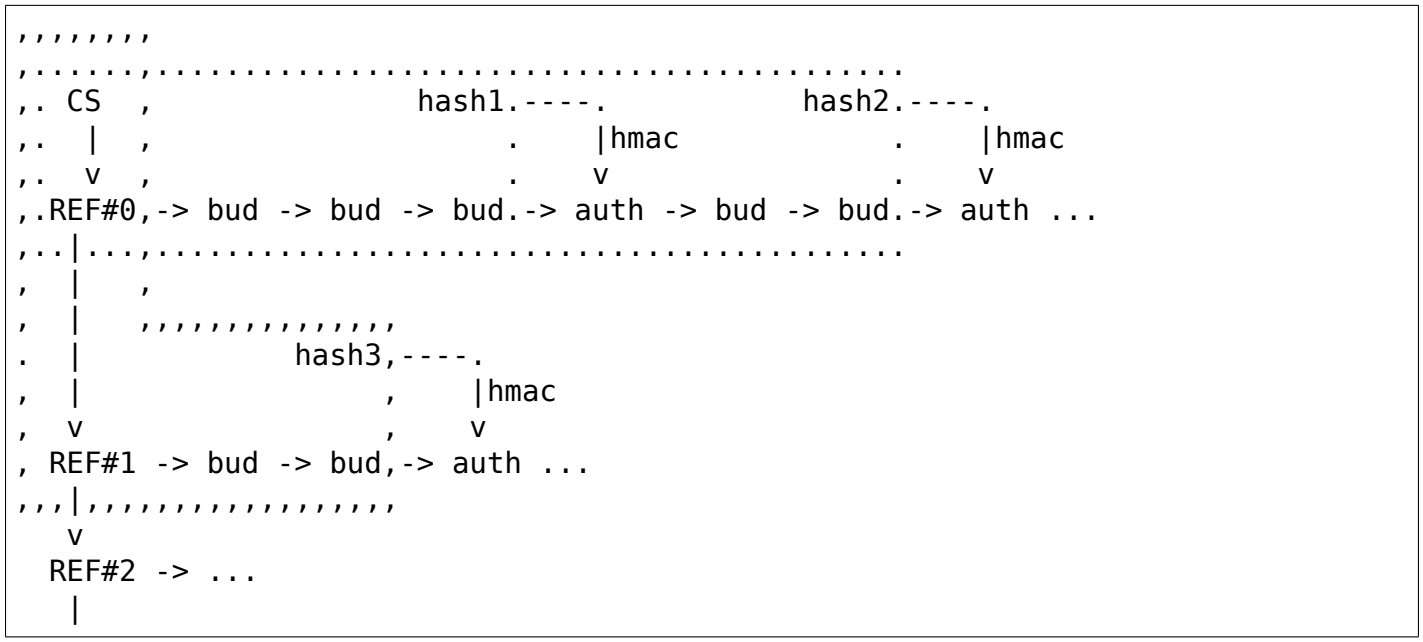

Figure 3: Coverage areas of index node hash and master node HMAC

The most important part for robustness and power-cut safety is to atomically persist the hash and file contents. Here the existing UBIFS logic for how changed nodes are persisted is already designed for this purpose such that UBIFS can safely recover if a power-cut occurs while persisting. Adding hashes to index nodes does not change this since each hash will be persisted atomically together with its respective node.

### Journal Authentication

The journal is authenticated too. Since the journal is continuously written it is necessary to also add authentication information frequently to the journal so that in case of a powercut not too much data can't be authenticated. This is done by creating a continuous hash beginning from the commit start node over the previous reference nodes, the current reference node, and the bud nodes. From time to time whenever it is suitable authentication nodes are added between the bud nodes. This new node type contains a HMAC over the current state of the hash chain. That way a journal can be authenticated up to the last authentication node. The tail of the journal which may not have a authentication node cannot be authenticated and is skipped during journal replay.

We get this picture for journal authentication:



```
  V
 ...
```

Since the hash also includes the reference nodes an attacker cannot reorder or skip any journal heads for replay. An attacker can only remove bud nodes or reference nodes from the end of the journal, effectively rewinding the filesystem at maximum back to the last commit.

The location of the log area is stored in the master node. Since the master node is authenticated with a HMAC as described above, it is not possible to tamper with that without detection. The size of the log area is specified when the filesystem is created using *mkfs.ubifs* and stored in the superblock node. To avoid tampering with this and other values stored there, a HMAC is added to the superblock struct. The superblock node is stored in LEB 0 and is only modified on feature flag or similar changes, but never on file changes.

### LPT Authentication

The location of the LPT root node on the flash is stored in the UBIFS master node. Since the LPT is written and read atomically on every commit, there is no need to authenticate individual nodes of the tree. It suffices to protect the integrity of the full LPT by a simple hash stored in the master node. Since the master node itself is authenticated, the LPTs authenticity can be verified by verifying the authenticity of the master node and comparing the LTP hash stored there with the hash computed from the read on-flash LPT.

### Key Management

For simplicity, UBIFS authentication uses a single key to compute the HMACs of superblock, master, commit start and reference nodes. This key has to be available on creation of the filesystem (*mkfs.ubifs*) to authenticate the superblock node. Further, it has to be available on mount of the filesystem to verify authenticated nodes and generate new HMACs for changes.

UBIFS authentication is intended to operate side-by-side with UBIFS encryption (fscrypt) to provide confidentiality and authenticity. Since UBIFS encryption has a different approach of encryption policies per directory, there can be multiple fscrypt master keys and there might be folders without encryption. UBIFS authentication on the other hand has an all-or-nothing approach in the sense that it either authenticates everything of the filesystem or nothing. Because of this and because UBIFS authentication should also be usable without encryption, it does not share the same master key with fscrypt, but manages a dedicated authentication key.

The API for providing the authentication key has yet to be defined, but the key can eg. be provided by userspace through a keyring similar to the way it is currently done in fscrypt. It should however be noted that the current fscrypt approach has shown its flaws and the userspace API will eventually change [FSCRYPT-POLICY2].

Nevertheless, it will be possible for a user to provide a single passphrase or key in userspace that covers UBIFS authentication and encryption. This can be solved by the corresponding userspace tools which derive a second key for authentication in addition to the derived fscrypt master key used for encryption.

To be able to check if the proper key is available on mount, the UBIFS superblock node will additionally store a hash of the authentication key. This approach is similar to the approach proposed for fscrypt encryption policy v2 [FSCRYPT-POLICY2].

### 3.55.3 Future Extensions

In certain cases where a vendor wants to provide an authenticated filesystem image to customers, it should be possible to do so without sharing the secret UBIFS authentication key. Instead, in addition the each HMAC a digital signature could be stored where the vendor shares the public key alongside the filesystem image. In case this filesystem has to be modified afterwards, UBIFS can exchange all digital signatures with HMACs on first mount similar to the way the IMA/EVM subsystem deals with such situations. The HMAC key will then have to be provided beforehand in the normal way.

### 3.55.4 References

[CRYPTSETUP2] https://www.saout.de/pipermail/dm-crypt/2017-November/005745.html

[DMC-CBC-ATTACK] https://www.jakoblell.com/blog/2013/12/22/practical-malleability-attack-against

[DM-INTEGRITY] https://www.kernel.org/doc/Documentation/device-mapper/dm-integrity.rst

[DM-VERITY] https://www.kernel.org/doc/Documentation/device-mapper/verity.rst

[FSCRYPT-POLICY2] https://www.spinics.net/lists/linux-ext4/msg58710.html

[UBIFS-WP] http://www.linux-mtd.infradead.org/doc/ubifs_whitepaper.pdf

## 3.56 UDF file system

If you encounter problems with reading UDF discs using this driver, please report them according to MAINTAINERS file.

Write support requires a block driver which supports writing. Currently dvd+rw drives and media support true random sector writes, and so a udf filesystem on such devices can be directly mounted read/write. CD-RW media however, does not support this. Instead the media can be formatted for packet mode using the utility cdrwtool, then the pktcdvd driver can be bound to the underlying cd device to provide the required buffering and read-modify-write cycles to allow the filesystem random sector writes while providing the hardware with only full packet writes. While not required for dvd+rw media, use of the pktcdvd driver often enhances performance due to very poor read-modify-write support supplied internally by drive firmware.

---

The following mount options are supported:

| | |
|---|---|
| gid= | Set the default group. |
| umask= | Set the default umask. |
| mode= | Set the default file permissions. |
| dmode= | Set the default directory permissions. |
| uid= | Set the default user. |
| bs= | Set the block size. |
| unhide | Show otherwise hidden files. |
| undelete | Show deleted files in lists. |
| adinicb | Embed data in the inode (default) |
| noadinicb | Don't embed data in the inode |
| shortad | Use short ad's |
| longad | Use long ad's (default) |
| nostrict | Unset strict conformance |
| iocharset= | Set the NLS character set |

The uid= and gid= options need a bit more explaining. They will accept a decimal numeric value and all inodes on that mount will then appear as belonging to that uid and gid. Mount options also accept the string ”forget”. The forget option causes all IDs to be written to disk as -1 which is a way of UDF standard to indicate that IDs are not supported for these files .

For typical desktop use of removable media, you should set the ID to that of the interactively logged on user, and also specify the forget option. This way the interactive user will always see the files on the disk as belonging to him.

The remaining are for debugging and disaster recovery:

| | |
|---|---|
| novrs | Skip volume sequence recognition |

The following expect a offset from 0.

| | |
|---|---|
| session= | Set the CDROM session (default= last session) |
| anchor= | Override standard anchor location. (default= 256) |
| lastblock= | Set the last block of the filesystem/ |

---

**For the latest version and toolset see:**

https://github.com/pali/udftools

**Documentation on UDF and ECMA 167 is available FREE from:**

- http://www.osta.org/
- https://www.ecma-international.org/

## 3.57 virtiofs: virtio-fs host<->guest shared file system

- Copyright (C) 2019 Red Hat, Inc.

### 3.57.1 Introduction

The virtiofs file system for Linux implements a driver for the paravirtualized VIRTIO "virtio-fs" device for guest<->host file system sharing. It allows a guest to mount a directory that has been exported on the host.

Guests often require access to files residing on the host or remote systems. Use cases include making files available to new guests during installation, booting from a root file system located on the host, persistent storage for stateless or ephemeral guests, and sharing a directory between guests.

Although it is possible to use existing network file systems for some of these tasks, they require configuration steps that are hard to automate and they expose the storage network to the guest. The virtio-fs device was designed to solve these problems by providing file system access without networking.

Furthermore the virtio-fs device takes advantage of the co-location of the guest and host to increase performance and provide semantics that are not possible with network file systems.

### 3.57.2 Usage

Mount file system with tag `myfs` on `/mnt`:

```
guest# mount -t virtiofs myfs /mnt
```

Please see https://virtio-fs.gitlab.io/ for details on how to configure QEMU and the virtiofsd daemon.

#### Mount options

virtiofs supports general VFS mount options, for example, remount, ro, rw, context, etc. It also supports FUSE mount options.

#### atime behavior

The atime-related mount options, for example, noatime, strictatime, are ignored. The atime behavior for virtiofs is the same as the underlying filesystem of the directory that has been exported on the host.

### 3.57.3 Internals

Since the virtio-fs device uses the FUSE protocol for file system requests, the virtiofs file system for Linux is integrated closely with the FUSE file system client. The guest acts as the FUSE client while the host acts as the FUSE server. The /dev/fuse interface between the kernel and userspace is replaced with the virtio-fs device interface.

FUSE requests are placed into a virtqueue and processed by the host. The response portion of the buffer is filled in by the host and the guest handles the request completion.

Mapping /dev/fuse to virtqueues requires solving differences in semantics between /dev/fuse and virtqueues. Each time the /dev/fuse device is read, the FUSE client may choose which request to transfer, making it possible to prioritize certain requests over others. Virtqueues have queue semantics and it is not possible to change the order of requests that have been enqueued. This is especially important if the virtqueue becomes full since it is then impossible to add high priority requests. In order to address this difference, the virtio-fs device uses a ”hiprio” virtqueue specifically for requests that have priority over normal requests.

## 3.58 VFAT

### 3.58.1 USING VFAT

To use the vfat filesystem, use the filesystem type 'vfat'. i.e.:

```
mount -t vfat /dev/fd0 /mnt
```

No special partition formatter is required, 'mkdosfs' will work fine if you want to format from within Linux.

### 3.58.2 VFAT MOUNT OPTIONS

**uid=###**
Set the owner of all files on this filesystem. The default is the uid of current process.

**gid=###**
Set the group of all files on this filesystem. The default is the gid of current process.

**umask=###**
The permission mask (for files and directories, see *umask(1)*). The default is the umask of current process.

**dmask=###**
The permission mask for the directory. The default is the umask of current process.

**fmask=###**
The permission mask for files. The default is the umask of current process.

**allow_utime=###**
This option controls the permission check of mtime/atime.

**-20**: If current process is in group of file's group ID, you can change timestamp.

**-2**: Other users can change timestamp.

The default is set from dmask option. If the directory is writable, utime(2) is also allowed. i.e. ~dmask & 022.

Normally utime(2) checks current process is owner of the file, or it has CAP_FOWNER capability. But FAT filesystem doesn't have uid/gid on disk, so normal check is too inflexible. With this option you can relax it.

**codepage=###**
Sets the codepage number for converting to shortname characters on FAT filesystem. By default, FAT_DEFAULT_CODEPAGE setting is used.

**iocharset=<name>**
Character set to use for converting between the encoding is used for user visible filename and 16 bit Unicode characters. Long filenames are stored on disk in Unicode format, but Unix for the most part doesn't know how to deal with Unicode. By default, FAT_DEFAULT_IOCHARSET setting is used.

There is also an option of doing UTF-8 translations with the utf8 option.

---

**Note:** `iocharset=utf8` is not recommended. If unsure, you should consider the utf8 option instead.

---

**utf8=<bool>**
UTF-8 is the filesystem safe version of Unicode that is used by the console. It can be enabled or disabled for the filesystem with this option. If 'uni_xlate' gets set, UTF-8 gets disabled. By default, FAT_DEFAULT_UTF8 setting is used.

**uni_xlate=<bool>**
Translate unhandled Unicode characters to special escaped sequences. This would let you backup and restore filenames that are created with any Unicode characters. Until Linux supports Unicode for real, this gives you an alternative. Without this option, a '?' is used when no translation is possible. The escape character is ':' because it is otherwise illegal on the vfat filesystem. The escape sequence that gets used is ':' and the four digits of hexadecimal unicode.

**nonumtail=<bool>**
When creating 8.3 aliases, normally the alias will end in '~1' or tilde followed by some number. If this option is set, then if the filename is ”longfilename.txt” and ”longfile.txt” does not currently exist in the directory, longfile.txt will be the short alias instead of longfi~1.txt.

**usefree**
Use the ”free clusters” value stored on FSINFO. It will be used to determine number of free clusters without scanning disk. But it's not used by default, because recent Windows don't update it correctly in some case. If you are sure the ”free clusters” on FSINFO is correct, by this option you can avoid scanning disk.

**quiet**
Stops printing certain warning messages.

**check=s|r|n**
Case sensitivity checking setting.

**s**: strict, case sensitive

**r**: relaxed, case insensitive

**n**: normal, default setting, currently case insensitive

**nocase**
This was deprecated for vfat. Use `shortname=win95` instead.

**shortname=lower|win95|winnt|mixed**
Shortname display/create setting.

**lower**: convert to lowercase for display, emulate the Windows 95 rule for create.

**win95**: emulate the Windows 95 rule for display/create.

**winnt**: emulate the Windows NT rule for display/create.

**mixed**: emulate the Windows NT rule for display, emulate the Windows 95 rule for create.

Default setting is *mixed*.

**tz=UTC**
Interpret timestamps as UTC rather than local time. This option disables the conversion of timestamps between local time (as used by Windows on FAT) and UTC (which Linux uses internally). This is particularly useful when mounting devices (like digital cameras) that are set to UTC in order to avoid the pitfalls of local time.

**time_offset=minutes**
Set offset for conversion of timestamps from local time used by FAT to UTC. I.e. <minutes> minutes will be subtracted from each timestamp to convert it to UTC used internally by Linux. This is useful when time zone set in `sys_tz` is not the time zone used by the filesystem. Note that this option still does not provide correct time stamps in all cases in presence of DST - time stamps in a different DST setting will be off by one hour.

**showexec**
If set, the execute permission bits of the file will be allowed only if the extension part of the name is .EXE, .COM, or .BAT. Not set by default.

**debug**
Can be set, but unused by the current implementation.

**sys_immutable**
If set, ATTR_SYS attribute on FAT is handled as IMMUTABLE flag on Linux. Not set by default.

**flush**
If set, the filesystem will try to flush to disk more early than normal. Not set by default.

**rodir**
FAT has the ATTR_RO (read-only) attribute. On Windows, the ATTR_RO of the directory will just be ignored, and is used only by applications as a flag (e.g. it's set for the customized folder).

If you want to use ATTR_RO as read-only flag even for the directory, set this option.

**errors=panic|continue|remount-ro**
specify FAT behavior on critical errors: panic, continue without doing anything or remount the partition in read-only mode (default behavior).

**discard**
If set, issues discard/TRIM commands to the block device when blocks are freed. This is useful for SSD devices and sparse/thinly-provisioned LUNs.

**nfs=stale_rw|nostale_ro**
Enable this only if you want to export the FAT filesystem over NFS.

**stale_rw**: This option maintains an index (cache) of directory *inodes* by *i_logstart* which is used by the nfs-related code to improve look-ups. Full file operations (read/write) over NFS is supported but with cache eviction at NFS server, this could result in ESTALE issues.

**nostale_ro**: This option bases the *inode* number and filehandle on the on-disk location of a file in the MS-DOS directory entry. This ensures that ESTALE will not be returned after a file is evicted from the inode cache. However, it means that operations such as rename, create and unlink could cause filehandles that previously pointed at one file to point at a different file, potentially causing data corruption. For this reason, this option also mounts the filesystem readonly.

To maintain backward compatibility, `'-o nfs'` is also accepted, defaulting to ”stale_rw”.

**dos1xfloppy <bool>: 0,1,yes,no,true,false**
If set, use a fallback default BIOS Parameter Block configuration, determined by backing device size. These static parameters match defaults assumed by DOS 1.x for 160 kiB, 180 kiB, 320 kiB, and 360 kiB floppies and floppy images.

### 3.58.3 LIMITATION

The fallocated region of file is discarded at umount/evict time when using fallocate with FALLOC_FL_KEEP_SIZE. So, User should assume that fallocated region can be discarded at last close if there is memory pressure resulting in eviction of the inode from the memory. As a result, for any dependency on the fallocated region, user should make sure to recheck fallocate after reopening the file.

### 3.58.4 TODO

Need to get rid of the raw scanning stuff. Instead, always use a get next directory entry approach. The only thing left that uses raw scanning is the directory renaming code.

### 3.58.5 POSSIBLE PROBLEMS

- vfat_valid_longname does not properly checked reserved names.
- When a volume name is the same as a directory name in the root directory of the filesystem, the directory name sometimes shows up as an empty file.
- autoconv option does not work correctly.

### 3.58.6 TEST SUITE

If you plan to make any modifications to the vfat filesystem, please get the test suite that comes with the vfat distribution at

http://web.archive.org/web/*/http://bmrc.berkeley.edu/people/chaffee/vfat.html

This tests quite a few parts of the vfat filesystem and additional tests for new features or untested features would be appreciated.

### 3.58.7 NOTES ON THE STRUCTURE OF THE VFAT FILESYSTEM

This documentation was provided by Galen C. Hunt gchunt@cs.rochester.edu and lightly annotated by Gordon Chaffee.

This document presents a very rough, technical overview of my knowledge of the extended FAT file system used in Windows NT 3.5 and Windows 95. I don't guarantee that any of the following is correct, but it appears to be so.

The extended FAT file system is almost identical to the FAT file system used in DOS versions up to and including *6.223410239847* :-). The significant change has been the addition of long file names. These names support up to 255 characters including spaces and lower case characters as opposed to the traditional 8.3 short names.

Here is the description of the traditional FAT entry in the current Windows 95 filesystem:

```
struct directory { // Short 8.3 names
        unsigned char name[8];          // file name
        unsigned char ext[3];           // file extension
        unsigned char attr;             // attribute byte
        unsigned char lcase;            // Case for base and extension
        unsigned char ctime_ms;         // Creation time, milliseconds
        unsigned char ctime[2];         // Creation time
        unsigned char cdate[2];         // Creation date
        unsigned char adate[2];         // Last access date
        unsigned char reserved[2];      // reserved values (ignored)
        unsigned char time[2];          // time stamp
        unsigned char date[2];          // date stamp
        unsigned char start[2];         // starting cluster number
        unsigned char size[4];          // size of the file
};
```

The lcase field specifies if the base and/or the extension of an 8.3 name should be capitalized. This field does not seem to be used by Windows 95 but it is used by Windows NT. The case of filenames is not completely compatible from Windows NT to Windows 95. It is not completely compatible in the reverse direction, however. Filenames that fit in the 8.3 namespace and are written on Windows NT to be lowercase will show up as uppercase on Windows 95.

---

**Note:** Note that the `start` and `size` values are actually little endian integer values. The descriptions of the fields in this structure are public knowledge and can be found elsewhere.

---

With the extended FAT system, Microsoft has inserted extra directory entries for any files with extended names. (Any name which legally fits within the old 8.3 encoding scheme does not have extra entries.) I call these extra entries slots. Basically, a slot is a specially formatted directory entry which holds up to 13 characters of a file's extended name. Think of slots as additional labeling for the directory entry of the file to which they correspond. Microsoft prefers to refer to the 8.3 entry for a file as its alias and the extended slot directory entries as the file name.

The C structure for a slot directory entry follows:

```
struct slot { // Up to 13 characters of a long name
        unsigned char id;               // sequence number for slot
        unsigned char name0_4[10];      // first 5 characters in name
        unsigned char attr;             // attribute byte
        unsigned char reserved;         // always 0
        unsigned char alias_checksum;   // checksum for 8.3 alias
        unsigned char name5_10[12];     // 6 more characters in name
        unsigned char start[2];         // starting cluster number
        unsigned char name11_12[4];     // last 2 characters in name
};
```

If the layout of the slots looks a little odd, it's only because of Microsoft's efforts to maintain compatibility with old software. The slots must be disguised to prevent old software from panicking. To this end, a number of measures are taken:

1) The attribute byte for a slot directory entry is always set to 0x0f. This corresponds to an old directory entry with attributes of ”hidden”, ”system”, ”read-only”, and ”volume label”. Most old software will ignore any directory entries with the ”volume label” bit set. Real volume label entries don't have the other three bits set.

2) The starting cluster is always set to 0, an impossible value for a DOS file.

Because the extended FAT system is backward compatible, it is possible for old software to modify directory entries. Measures must be taken to ensure the validity of slots. An extended FAT system can verify that a slot does in fact belong to an 8.3 directory entry by the following:

1) Positioning. Slots for a file always immediately proceed their corresponding 8.3 directory entry. In addition, each slot has an id which marks its order in the extended file name. Here is a very abbreviated view of an 8.3 directory entry and its corresponding long name slots for the file ”My Big File.Extension which is long”:

```
<proceeding files...>
<slot #3, id = 0x43, characters = "h is long">
<slot #2, id = 0x02, characters = "xtension whic">
<slot #1, id = 0x01, characters = "My Big File.E">
<directory entry, name = "MYBIGFIL.EXT">
```

---

**Note:** Note that the slots are stored from last to first. Slots are numbered from 1 to N. The Nth slot is `or'ed` with 0x40 to mark it as the last one.

---

2) Checksum. Each slot has an alias_checksum value. The checksum is calculated from the 8.3 name using the following algorithm:

```
for (sum = i = 0; i < 11; i++) {
        sum = (((sum&1)<<7)|((sum&0xfe)>>1)) + name[i]
}
```

3) If there is free space in the final slot, a Unicode `NULL (0x0000)` is stored after the final character. After that, all unused characters in the final slot are set to Unicode 0xFFFF.

Finally, note that the extended name is stored in Unicode. Each Unicode character takes either two or four bytes, UTF-16LE encoded.

## 3.59 XFS Filesystem Documentation

### 3.59.1 XFS Logging Design

#### Preamble

This document describes the design and algorithms that the XFS journalling subsystem is based on. This document describes the design and algorithms that the XFS journalling subsystem is based on so that readers may familiarize themselves with the general concepts of how transaction processing in XFS works.

We begin with an overview of transactions in XFS, followed by describing how transaction reservations are structured and accounted, and then move into how we guarantee forwards progress for long running transactions with finite initial reservations bounds. At this point we need to explain how relogging works. With the basic concepts covered, the design of the delayed logging mechanism is documented.

#### Introduction

XFS uses Write Ahead Logging for ensuring changes to the filesystem metadata are atomic and recoverable. For reasons of space and time efficiency, the logging mechanisms are varied and complex, combining intents, logical and physical logging mechanisms to provide the necessary recovery guarantees the filesystem requires.

Some objects, such as inodes and dquots, are logged in logical format where the details logged are made up of the changes to in-core structures rather than on-disk structures. Other objects - typically buffers - have their physical changes logged. Long running atomic modifications have individual changes chained together by intents, ensuring that journal recovery can restart and finish an operation that was only partially done when the system stopped functioning.

The reason for these differences is to keep the amount of log space and CPU time required to process objects being modified as small as possible and hence the logging overhead as low as possible. Some items are very frequently modified, and some parts of objects are more frequently modified than others, so keeping the overhead of metadata logging low is of prime importance.

The method used to log an item or chain modifications together isn't particularly important in the scope of this document. It suffices to know that the method used for logging a particular object or chaining modifications together are different and are dependent on the object and/or modification being performed. The logging subsystem only cares that certain specific rules are followed to guarantee forwards progress and prevent deadlocks.

### Transactions in XFS

XFS has two types of high level transactions, defined by the type of log space reservation they take. These are known as ”one shot” and ”permanent” transactions. Permanent transaction reservations can take reservations that span commit boundaries, whilst ”one shot” transactions are for a single atomic modification.

The type and size of reservation must be matched to the modification taking place. This means that permanent transactions can be used for one-shot modifications, but one-shot reservations cannot be used for permanent transactions.

In the code, a one-shot transaction pattern looks somewhat like this:

```
tp = xfs_trans_alloc(<reservation>)
<lock items>
<join item to transaction>
<do modification>
xfs_trans_commit(tp);
```

As items are modified in the transaction, the dirty regions in those items are tracked via the transaction handle. Once the transaction is committed, all resources joined to it are released, along with the remaining unused reservation space that was taken at the transaction allocation time.

In contrast, a permanent transaction is made up of multiple linked individual transactions, and the pattern looks like this:

```
tp = xfs_trans_alloc(<reservation>)
xfs_ilock(ip, XFS_ILOCK_EXCL)

loop {
        xfs_trans_ijoin(tp, 0);
        <do modification>
        xfs_trans_log_inode(tp, ip);
        xfs_trans_roll(&tp);
}

xfs_trans_commit(tp);
xfs_iunlock(ip, XFS_ILOCK_EXCL);
```

While this might look similar to a one-shot transaction, there is an important difference: xfs_trans_roll() performs a specific operation that links two transactions together:

```
ntp = xfs_trans_dup(tp);
xfs_trans_commit(tp);
xfs_trans_reserve(ntp);
```

This results in a series of ”rolling transactions” where the inode is locked across the entire chain of transactions. Hence while this series of rolling transactions is running, nothing else can read from or write to the inode and this provides a mechanism for complex changes to appear atomic from an external observer's point of view.

It is important to note that a series of rolling transactions in a permanent transaction does not form an atomic change in the journal. While each individual modification is atomic, the

chain is *not atomic*. If we crash half way through, then recovery will only replay up to the last transactional modification the loop made that was committed to the journal.

This affects long running permanent transactions in that it is not possible to predict how much of a long running operation will actually be recovered because there is no guarantee of how much of the operation reached stale storage. Hence if a long running operation requires multiple transactions to fully complete, the high level operation must use intents and deferred operations to guarantee recovery can complete the operation once the first transactions is persisted in the on-disk journal.

### Transactions are Asynchronous

In XFS, all high level transactions are asynchronous by default. This means that xfs_trans_commit() does not guarantee that the modification has been committed to stable storage when it returns. Hence when a system crashes, not all the completed transactions will be replayed during recovery.

However, the logging subsystem does provide global ordering guarantees, such that if a specific change is seen after recovery, all metadata modifications that were committed prior to that change will also be seen.

For single shot operations that need to reach stable storage immediately, or ensuring that a long running permanent transaction is fully committed once it is complete, we can explicitly tag a transaction as synchronous. This will trigger a ”log force” to flush the outstanding committed transactions to stable storage in the journal and wait for that to complete.

Synchronous transactions are rarely used, however, because they limit logging throughput to the IO latency limitations of the underlying storage. Instead, we tend to use log forces to ensure modifications are on stable storage only when a user operation requires a synchronisation point to occur (e.g. fsync).

### Transaction Reservations

It has been mentioned a number of times now that the logging subsystem needs to provide a forwards progress guarantee so that no modification ever stalls because it can't be written to the journal due to a lack of space in the journal. This is achieved by the transaction reservations that are made when a transaction is first allocated. For permanent transactions, these reservations are maintained as part of the transaction rolling mechanism.

A transaction reservation provides a guarantee that there is physical log space available to write the modification into the journal before we start making modifications to objects and items. As such, the reservation needs to be large enough to take into account the amount of metadata that the change might need to log in the worst case. This means that if we are modifying a btree in the transaction, we have to reserve enough space to record a full leaf-to-root split of the btree. As such, the reservations are quite complex because we have to take into account all the hidden changes that might occur.

For example, a user data extent allocation involves allocating an extent from free space, which modifies the free space trees. That's two btrees. Inserting the extent into the inode's extent map might require a split of the extent map btree, which requires another allocation that can modify the free space trees again. Then we might have to update reverse mappings, which modifies yet another btree which might require more space. And so on. Hence the amount of metadata that a ”simple” operation can modify can be quite large.

This ”worst case” calculation provides us with the static ”unit reservation” for the transaction that is calculated at mount time. We must guarantee that the log has this much space available before the transaction is allowed to proceed so that when we come to write the dirty metadata into the log we don't run out of log space half way through the write.

For one-shot transactions, a single unit space reservation is all that is required for the transaction to proceed. For permanent transactions, however, we also have a ”log count” that affects the size of the reservation that is to be made.

While a permanent transaction can get by with a single unit of space reservation, it is somewhat inefficient to do this as it requires the transaction rolling mechanism to re-reserve space on every transaction roll. We know from the implementation of the permanent transactions how many transaction rolls are likely for the common modifications that need to be made.

For example, an inode allocation is typically two transactions - one to physically allocate a free inode chunk on disk, and another to allocate an inode from an inode chunk that has free inodes in it. Hence for an inode allocation transaction, we might set the reservation log count to a value of 2 to indicate that the common/fast path transaction will commit two linked transactions in a chain. Each time a permanent transaction rolls, it consumes an entire unit reservation.

Hence when the permanent transaction is first allocated, the log space reservation is increased from a single unit reservation to multiple unit reservations. That multiple is defined by the reservation log count, and this means we can roll the transaction multiple times before we have to re-reserve log space when we roll the transaction. This ensures that the common modifications we make only need to reserve log space once.

If the log count for a permanent transaction reaches zero, then it needs to re-reserve physical space in the log. This is somewhat complex, and requires an understanding of how the log accounts for space that has been reserved.

### Log Space Accounting

The position in the log is typically referred to as a Log Sequence Number (LSN). The log is circular, so the positions in the log are defined by the combination of a cycle number - the number of times the log has been overwritten - and the offset into the log. A LSN carries the cycle in the upper 32 bits and the offset in the lower 32 bits. The offset is in units of ”basic blocks” (512 bytes). Hence we can do realtively simple LSN based math to keep track of available space in the log.

Log space accounting is done via a pair of constructs called ”grant heads”. The position of the grant heads is an absolute value, so the amount of space available in the log is defined by the distance between the position of the grant head and the current log tail. That is, how much space can be reserved/consumed before the grant heads would fully wrap the log and overtake the tail position.

The first grant head is the ”reserve” head. This tracks the byte count of the reservations currently held by active transactions. It is a purely in-memory accounting of the space reservation and, as such, actually tracks byte offsets into the log rather than basic blocks. Hence it technically isn't using LSNs to represent the log position, but it is still treated like a split {cycle,offset} tuple for the purposes of tracking reservation space.

The reserve grant head is used to accurately account for exact transaction reservations amounts and the exact byte count that modifications actually make and need to write into the log. The reserve head is used to prevent new transactions from taking new reservations when the head reaches the current tail. It will block new reservations in a FIFO queue and as the log tail moves

forward it will wake them in order once sufficient space is available. This FIFO mechanism ensures no transaction is starved of resources when log space shortages occur.

The other grant head is the ”write” head. Unlike the reserve head, this grant head contains an LSN and it tracks the physical space usage in the log. While this might sound like it is accounting the same state as the reserve grant head - and it mostly does track exactly the same location as the reserve grant head - there are critical differences in behaviour between them that provides the forwards progress guarantees that rolling permanent transactions require.

These differences when a permanent transaction is rolled and the internal ”log count” reaches zero and the initial set of unit reservations have been exhausted. At this point, we still require a log space reservation to continue the next transaction in the sequeunce, but we have none remaining. We cannot sleep during the transaction commit process waiting for new log space to become available, as we may end up on the end of the FIFO queue and the items we have locked while we sleep could end up pinning the tail of the log before there is enough free space in the log to fulfill all of the pending reservations and then wake up transaction commit in progress.

To take a new reservation without sleeping requires us to be able to take a reservation even if there is no reservation space currently available. That is, we need to be able to *overcommit* the log reservation space. As has already been detailed, we cannot overcommit physical log space. However, the reserve grant head does not track physical space - it only accounts for the amount of reservations we currently have outstanding. Hence if the reserve head passes over the tail of the log all it means is that new reservations will be throttled immediately and remain throttled until the log tail is moved forward far enough to remove the overcommit and start taking new reservations. In other words, we can overcommit the reserve head without violating the physical log head and tail rules.

As a result, permanent transactions only ”regrant” reservation space during xfs_trans_commit() calls, while the physical log space reservation - tracked by the write head - is then reserved separately by a call to xfs_log_reserve() after the commit completes. Once the commit completes, we can sleep waiting for physical log space to be reserved from the write grant head, but only if one critical rule has been observed:

```
Code using permanent reservations must always log the items they hold
locked across each transaction they roll in the chain.
```

”Re-logging” the locked items on every transaction roll ensures that the items attached to the transaction chain being rolled are always relocated to the physical head of the log and so do not pin the tail of the log. If a locked item pins the tail of the log when we sleep on the write reservation, then we will deadlock the log as we cannot take the locks needed to write back that item and move the tail of the log forwards to free up write grant space. Re-logging the locked items avoids this deadlock and guarantees that the log reservation we are making cannot self-deadlock.

If all rolling transactions obey this rule, then they can all make forwards progress independently because nothing will block the progress of the log tail moving forwards and hence ensuring that write grant space is always (eventually) made available to permanent transactions no matter how many times they roll.

### Re-logging Explained

XFS allows multiple separate modifications to a single object to be carried in the log at any given time. This allows the log to avoid needing to flush each change to disk before recording a new change to the object. XFS does this via a method called ”re-logging”. Conceptually, this is quite simple - all it requires is that any new change to the object is recorded with a *new copy* of all the existing changes in the new transaction that is written to the log.

That is, if we have a sequence of changes A through to F, and the object was written to disk after change D, we would see in the log the following series of transactions, their contents and the log sequence number (LSN) of the transaction:

```
Transaction             Contents        LSN
   A                       A               X
   B                      A+B             X+n
   C                     A+B+C           X+n+m
   D                    A+B+C+D         X+n+m+o
    <object written to disk>
   E                       E               Y (> X+n+m+o)
   F                      E+F             Y+p
```

In other words, each time an object is relogged, the new transaction contains the aggregation of all the previous changes currently held only in the log.

This relogging technique allows objects to be moved forward in the log so that an object being relogged does not prevent the tail of the log from ever moving forward. This can be seen in the table above by the changing (increasing) LSN of each subsequent transaction, and it's the technique that allows us to implement long-running, multiple-commit permanent transactions.

A typical example of a rolling transaction is the removal of extents from an inode which can only be done at a rate of two extents per transaction because of reservation size limitations. Hence a rolling extent removal transaction keeps relogging the inode and btree buffers as they get modified in each removal operation. This keeps them moving forward in the log as the operation progresses, ensuring that current operation never gets blocked by itself if the log wraps around.

Hence it can be seen that the relogging operation is fundamental to the correct working of the XFS journalling subsystem. From the above description, most people should be able to see why the XFS metadata operations writes so much to the log - repeated operations to the same objects write the same changes to the log over and over again. Worse is the fact that objects tend to get dirtier as they get relogged, so each subsequent transaction is writing more metadata into the log.

It should now also be obvious how relogging and asynchronous transactions go hand in hand. That is, transactions don't get written to the physical journal until either a log buffer is filled (a log buffer can hold multiple transactions) or a synchronous operation forces the log buffers holding the transactions to disk. This means that XFS is doing aggregation of transactions in memory - batching them, if you like - to minimise the impact of the log IO on transaction throughput.

The limitation on asynchronous transaction throughput is the number and size of log buffers made available by the log manager. By default there are 8 log buffers available and the size of each is 32kB - the size can be increased up to 256kB by use of a mount option.

Effectively, this gives us the maximum bound of outstanding metadata changes that can be

made to the filesystem at any point in time - if all the log buffers are full and under IO, then no more transactions can be committed until the current batch completes. It is now common for a single current CPU core to be to able to issue enough transactions to keep the log buffers full and under IO permanently. Hence the XFS journalling subsystem can be considered to be IO bound.

### Delayed Logging: Concepts

The key thing to note about the asynchronous logging combined with the relogging technique XFS uses is that we can be relogging changed objects multiple times before they are committed to disk in the log buffers. If we return to the previous relogging example, it is entirely possible that transactions A through D are committed to disk in the same log buffer.

That is, a single log buffer may contain multiple copies of the same object, but only one of those copies needs to be there - the last one "D", as it contains all the changes from the previous changes. In other words, we have one necessary copy in the log buffer, and three stale copies that are simply wasting space. When we are doing repeated operations on the same set of objects, these "stale objects" can be over 90% of the space used in the log buffers. It is clear that reducing the number of stale objects written to the log would greatly reduce the amount of metadata we write to the log, and this is the fundamental goal of delayed logging.

From a conceptual point of view, XFS is already doing relogging in memory (where memory == log buffer), only it is doing it extremely inefficiently. It is using logical to physical formatting to do the relogging because there is no infrastructure to keep track of logical changes in memory prior to physically formatting the changes in a transaction to the log buffer. Hence we cannot avoid accumulating stale objects in the log buffers.

Delayed logging is the name we've given to keeping and tracking transactional changes to objects in memory outside the log buffer infrastructure. Because of the relogging concept fundamental to the XFS journalling subsystem, this is actually relatively easy to do - all the changes to logged items are already tracked in the current infrastructure. The big problem is how to accumulate them and get them to the log in a consistent, recoverable manner. Describing the problems and how they have been solved is the focus of this document.

One of the key changes that delayed logging makes to the operation of the journalling subsystem is that it disassociates the amount of outstanding metadata changes from the size and number of log buffers available. In other words, instead of there only being a maximum of 2MB of transaction changes not written to the log at any point in time, there may be a much greater amount being accumulated in memory. Hence the potential for loss of metadata on a crash is much greater than for the existing logging mechanism.

It should be noted that this does not change the guarantee that log recovery will result in a consistent filesystem. What it does mean is that as far as the recovered filesystem is concerned, there may be many thousands of transactions that simply did not occur as a result of the crash. This makes it even more important that applications that care about their data use fsync() where they need to ensure application level data integrity is maintained.

It should be noted that delayed logging is not an innovative new concept that warrants rigorous proofs to determine whether it is correct or not. The method of accumulating changes in memory for some period before writing them to the log is used effectively in many filesystems including ext3 and ext4. Hence no time is spent in this document trying to convince the reader that the concept is sound. Instead it is simply considered a "solved problem" and as such implementing it in XFS is purely an exercise in software engineering.

The fundamental requirements for delayed logging in XFS are simple:

1. Reduce the amount of metadata written to the log by at least an order of magnitude.
2. Supply sufficient statistics to validate Requirement #1.
3. Supply sufficient new tracing infrastructure to be able to debug problems with the new code.
4. No on-disk format change (metadata or log format).
5. Enable and disable with a mount option.
6. No performance regressions for synchronous transaction workloads.

### Delayed Logging: Design

#### Storing Changes

The problem with accumulating changes at a logical level (i.e. just using the existing log item dirty region tracking) is that when it comes to writing the changes to the log buffers, we need to ensure that the object we are formatting is not changing while we do this. This requires locking the object to prevent concurrent modification. Hence flushing the logical changes to the log would require us to lock every object, format them, and then unlock them again.

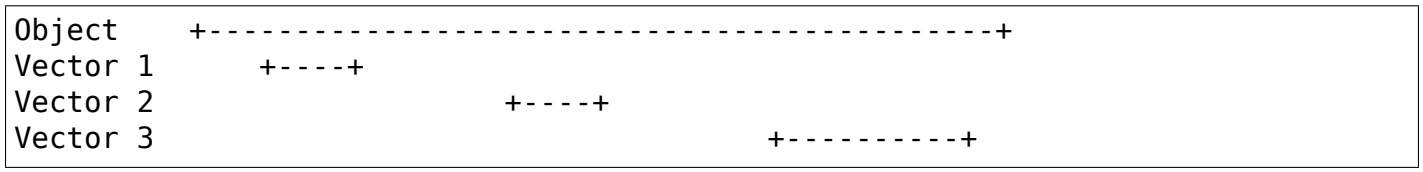
This introduces lots of scope for deadlocks with transactions that are already running. For example, a transaction has object A locked and modified, but needs the delayed logging tracking lock to commit the transaction. However, the flushing thread has the delayed logging tracking lock already held, and is trying to get the lock on object A to flush it to the log buffer. This appears to be an unsolvable deadlock condition, and it was solving this problem that was the barrier to implementing delayed logging for so long.

The solution is relatively simple - it just took a long time to recognise it. Put simply, the current logging code formats the changes to each item into an vector array that points to the changed regions in the item. The log write code simply copies the memory these vectors point to into the log buffer during transaction commit while the item is locked in the transaction. Instead of using the log buffer as the destination of the formatting code, we can use an allocated memory buffer big enough to fit the formatted vector.

If we then copy the vector into the memory buffer and rewrite the vector to point to the memory buffer rather than the object itself, we now have a copy of the changes in a format that is compatible with the log buffer writing code. that does not require us to lock the item to access. This formatting and rewriting can all be done while the object is locked during transaction commit, resulting in a vector that is transactionally consistent and can be accessed without needing to lock the owning item.

Hence we avoid the need to lock items when we need to flush outstanding asynchronous transactions to the log. The differences between the existing formatting method and the delayed logging formatting can be seen in the diagram below.

Current format log vector:



After formatting:

```
Log Buffer    +-V1-+-V2-+----V3----+
```

Delayed logging vector:

```
Object    +---------------------------------------------+
Vector 1      +----+
Vector 2                   +----+
Vector 3                                 +----------+
```

After formatting:

```
Memory Buffer +-V1-+-V2-+----V3----+
Vector 1      +----+
Vector 2           +----+
Vector 3                +----------+
```

The memory buffer and associated vector need to be passed as a single object, but still need to be associated with the parent object so if the object is relogged we can replace the current memory buffer with a new memory buffer that contains the latest changes.

The reason for keeping the vector around after we've formatted the memory buffer is to support splitting vectors across log buffer boundaries correctly. If we don't keep the vector around, we do not know where the region boundaries are in the item, so we'd need a new encapsulation method for regions in the log buffer writing (i.e. double encapsulation). This would be an on-disk format change and as such is not desirable. It also means we'd have to write the log region headers in the formatting stage, which is problematic as there is per region state that needs to be placed into the headers during the log write.

Hence we need to keep the vector, but by attaching the memory buffer to it and rewriting the vector addresses to point at the memory buffer we end up with a self-describing object that can be passed to the log buffer write code to be handled in exactly the same manner as the existing log vectors are handled. Hence we avoid needing a new on-disk format to handle items that have been relogged in memory.

### Tracking Changes

Now that we can record transactional changes in memory in a form that allows them to be used without limitations, we need to be able to track and accumulate them so that they can be written to the log at some later point in time. The log item is the natural place to store this vector and buffer, and also makes sense to be the object that is used to track committed objects as it will always exist once the object has been included in a transaction.

The log item is already used to track the log items that have been written to the log but not yet written to disk. Such log items are considered ”active” and as such are stored in the Active Item List (AIL) which is a LSN-ordered double linked list. Items are inserted into this list during log buffer IO completion, after which they are unpinned and can be written to disk. An object that is in the AIL can be relogged, which causes the object to be pinned again and then moved forward in the AIL when the log buffer IO completes for that transaction.

Essentially, this shows that an item that is in the AIL can still be modified and relogged, so any tracking must be separate to the AIL infrastructure. As such, we cannot reuse the AIL list

pointers for tracking committed items, nor can we store state in any field that is protected by the AIL lock. Hence the committed item tracking needs its own locks, lists and state fields in the log item.

Similar to the AIL, tracking of committed items is done through a new list called the Committed Item List (CIL). The list tracks log items that have been committed and have formatted memory buffers attached to them. It tracks objects in transaction commit order, so when an object is relogged it is removed from its place in the list and re-inserted at the tail. This is entirely arbitrary and done to make it easy for debugging - the last items in the list are the ones that are most recently modified. Ordering of the CIL is not necessary for transactional integrity (as discussed in the next section) so the ordering is done for convenience/sanity of the developers.

### Delayed Logging: Checkpoints

When we have a log synchronisation event, commonly known as a ”log force”, all the items in the CIL must be written into the log via the log buffers. We need to write these items in the order that they exist in the CIL, and they need to be written as an atomic transaction. The need for all the objects to be written as an atomic transaction comes from the requirements of relogging and log replay - all the changes in all the objects in a given transaction must either be completely replayed during log recovery, or not replayed at all. If a transaction is not replayed because it is not complete in the log, then no later transactions should be replayed, either.

To fulfill this requirement, we need to write the entire CIL in a single log transaction. Fortunately, the XFS log code has no fixed limit on the size of a transaction, nor does the log replay code. The only fundamental limit is that the transaction cannot be larger than just under half the size of the log. The reason for this limit is that to find the head and tail of the log, there must be at least one complete transaction in the log at any given time. If a transaction is larger than half the log, then there is the possibility that a crash during the write of a such a transaction could partially overwrite the only complete previous transaction in the log. This will result in a recovery failure and an inconsistent filesystem and hence we must enforce the maximum size of a checkpoint to be slightly less than a half the log.

Apart from this size requirement, a checkpoint transaction looks no different to any other transaction - it contains a transaction header, a series of formatted log items and a commit record at the tail. From a recovery perspective, the checkpoint transaction is also no different - just a lot bigger with a lot more items in it. The worst case effect of this is that we might need to tune the recovery transaction object hash size.

Because the checkpoint is just another transaction and all the changes to log items are stored as log vectors, we can use the existing log buffer writing code to write the changes into the log. To do this efficiently, we need to minimise the time we hold the CIL locked while writing the checkpoint transaction. The current log write code enables us to do this easily with the way it separates the writing of the transaction contents (the log vectors) from the transaction commit record, but tracking this requires us to have a per-checkpoint context that travels through the log write process through to checkpoint completion.

Hence a checkpoint has a context that tracks the state of the current checkpoint from initiation to checkpoint completion. A new context is initiated at the same time a checkpoint transaction is started. That is, when we remove all the current items from the CIL during a checkpoint operation, we move all those changes into the current checkpoint context. We then initialise a new context and attach that to the CIL for aggregation of new transactions.

This allows us to unlock the CIL immediately after transfer of all the committed items and

effectively allows new transactions to be issued while we are formatting the checkpoint into the log. It also allows concurrent checkpoints to be written into the log buffers in the case of log force heavy workloads, just like the existing transaction commit code does. This, however, requires that we strictly order the commit records in the log so that checkpoint sequence order is maintained during log replay.

To ensure that we can be writing an item into a checkpoint transaction at the same time another transaction modifies the item and inserts the log item into the new CIL, then checkpoint transaction commit code cannot use log items to store the list of log vectors that need to be written into the transaction. Hence log vectors need to be able to be chained together to allow them to be detached from the log items. That is, when the CIL is flushed the memory buffer and log vector attached to each log item needs to be attached to the checkpoint context so that the log item can be released. In diagrammatic form, the CIL would look like this before the flush:

```
CIL Head
   |
   V
Log Item <-> log vector 1       -> memory buffer
   |                            -> vector array
   V
Log Item <-> log vector 2       -> memory buffer
   |                            -> vector array
   V
......
   |
   V
Log Item <-> log vector N-1     -> memory buffer
   |                            -> vector array
   V
Log Item <-> log vector N       -> memory buffer
                                -> vector array
```

And after the flush the CIL head is empty, and the checkpoint context log vector list would look like:

```
Checkpoint Context
   |
   V
log vector 1    -> memory buffer
   |            -> vector array
   |            -> Log Item
   V
log vector 2    -> memory buffer
   |            -> vector array
   |            -> Log Item
   V
......
   |
   V
log vector N-1  -> memory buffer
   |            -> vector array
   |            -> Log Item
```

```
   V
log vector N    -> memory buffer
                -> vector array
                -> Log Item
```

Once this transfer is done, the CIL can be unlocked and new transactions can start, while the checkpoint flush code works over the log vector chain to commit the checkpoint.

Once the checkpoint is written into the log buffers, the checkpoint context is attached to the log buffer that the commit record was written to along with a completion callback. Log IO completion will call that callback, which can then run transaction committed processing for the log items (i.e. insert into AIL and unpin) in the log vector chain and then free the log vector chain and checkpoint context.

Discussion Point: I am uncertain as to whether the log item is the most efficient way to track vectors, even though it seems like the natural way to do it. The fact that we walk the log items (in the CIL) just to chain the log vectors and break the link between the log item and the log vector means that we take a cache line hit for the log item list modification, then another for the log vector chaining. If we track by the log vectors, then we only need to break the link between the log item and the log vector, which means we should dirty only the log item cachelines. Normally I wouldn't be concerned about one vs two dirty cachelines except for the fact I've seen upwards of 80,000 log vectors in one checkpoint transaction. I'd guess this is a ”measure and compare” situation that can be done after a working and reviewed implementation is in the dev tree....

### Delayed Logging: Checkpoint Sequencing

One of the key aspects of the XFS transaction subsystem is that it tags committed transactions with the log sequence number of the transaction commit. This allows transactions to be issued asynchronously even though there may be future operations that cannot be completed until that transaction is fully committed to the log. In the rare case that a dependent operation occurs (e.g. re-using a freed metadata extent for a data extent), a special, optimised log force can be issued to force the dependent transaction to disk immediately.

To do this, transactions need to record the LSN of the commit record of the transaction. This LSN comes directly from the log buffer the transaction is written into. While this works just fine for the existing transaction mechanism, it does not work for delayed logging because transactions are not written directly into the log buffers. Hence some other method of sequencing transactions is required.

As discussed in the checkpoint section, delayed logging uses per-checkpoint contexts, and as such it is simple to assign a sequence number to each checkpoint. Because the switching of checkpoint contexts must be done atomically, it is simple to ensure that each new context has a monotonically increasing sequence number assigned to it without the need for an external atomic counter - we can just take the current context sequence number and add one to it for the new context.

Then, instead of assigning a log buffer LSN to the transaction commit LSN during the commit, we can assign the current checkpoint sequence. This allows operations that track transactions that have not yet completed know what checkpoint sequence needs to be committed before they can continue. As a result, the code that forces the log to a specific LSN now needs to ensure that the log forces to a specific checkpoint.

To ensure that we can do this, we need to track all the checkpoint contexts that are currently committing to the log. When we flush a checkpoint, the context gets added to a ”committing” list which can be searched. When a checkpoint commit completes, it is removed from the committing list. Because the checkpoint context records the LSN of the commit record for the checkpoint, we can also wait on the log buffer that contains the commit record, thereby using the existing log force mechanisms to execute synchronous forces.

It should be noted that the synchronous forces may need to be extended with mitigation algorithms similar to the current log buffer code to allow aggregation of multiple synchronous transactions if there are already synchronous transactions being flushed. Investigation of the performance of the current design is needed before making any decisions here.

The main concern with log forces is to ensure that all the previous checkpoints are also committed to disk before the one we need to wait for. Therefore we need to check that all the prior contexts in the committing list are also complete before waiting on the one we need to complete. We do this synchronisation in the log force code so that we don't need to wait anywhere else for such serialisation - it only matters when we do a log force.

The only remaining complexity is that a log force now also has to handle the case where the forcing sequence number is the same as the current context. That is, we need to flush the CIL and potentially wait for it to complete. This is a simple addition to the existing log forcing code to check the sequence numbers and push if required. Indeed, placing the current sequence checkpoint flush in the log force code enables the current mechanism for issuing synchronous transactions to remain untouched (i.e. commit an asynchronous transaction, then force the log at the LSN of that transaction) and so the higher level code behaves the same regardless of whether delayed logging is being used or not.

### Delayed Logging: Checkpoint Log Space Accounting

The big issue for a checkpoint transaction is the log space reservation for the transaction. We don't know how big a checkpoint transaction is going to be ahead of time, nor how many log buffers it will take to write out, nor the number of split log vector regions are going to be used. We can track the amount of log space required as we add items to the commit item list, but we still need to reserve the space in the log for the checkpoint.

A typical transaction reserves enough space in the log for the worst case space usage of the transaction. The reservation accounts for log record headers, transaction and region headers, headers for split regions, buffer tail padding, etc. as well as the actual space for all the changed metadata in the transaction. While some of this is fixed overhead, much of it is dependent on the size of the transaction and the number of regions being logged (the number of log vectors in the transaction).

An example of the differences would be logging directory changes versus logging inode changes. If you modify lots of inode cores (e.g. `chmod -R g+w *`), then there are lots of transactions that only contain an inode core and an inode log format structure. That is, two vectors totaling roughly 150 bytes. If we modify 10,000 inodes, we have about 1.5MB of metadata to write in 20,000 vectors. Each vector is 12 bytes, so the total to be logged is approximately 1.75MB. In comparison, if we are logging full directory buffers, they are typically 4KB each, so we in 1.5MB of directory buffers we'd have roughly 400 buffers and a buffer format structure for each buffer - roughly 800 vectors or 1.51MB total space. From this, it should be obvious that a static log space reservation is not particularly flexible and is difficult to select the ”optimal value” for all workloads.

Further, if we are going to use a static reservation, which bit of the entire reservation does it cover? We account for space used by the transaction reservation by tracking the space currently used by the object in the CIL and then calculating the increase or decrease in space used as the object is relogged. This allows for a checkpoint reservation to only have to account for log buffer metadata used such as log header records.

However, even using a static reservation for just the log metadata is problematic. Typically log record headers use at least 16KB of log space per 1MB of log space consumed (512 bytes per 32k) and the reservation needs to be large enough to handle arbitrary sized checkpoint transactions. This reservation needs to be made before the checkpoint is started, and we need to be able to reserve the space without sleeping. For a 8MB checkpoint, we need a reservation of around 150KB, which is a non-trivial amount of space.

A static reservation needs to manipulate the log grant counters - we can take a permanent reservation on the space, but we still need to make sure we refresh the write reservation (the actual space available to the transaction) after every checkpoint transaction completion. Unfortunately, if this space is not available when required, then the regrant code will sleep waiting for it.

The problem with this is that it can lead to deadlocks as we may need to commit checkpoints to be able to free up log space (refer back to the description of rolling transactions for an example of this). Hence we *must* always have space available in the log if we are to use static reservations, and that is very difficult and complex to arrange. It is possible to do, but there is a simpler way.

The simpler way of doing this is tracking the entire log space used by the items in the CIL and using this to dynamically calculate the amount of log space required by the log metadata. If this log metadata space changes as a result of a transaction commit inserting a new memory buffer into the CIL, then the difference in space required is removed from the transaction that causes the change. Transactions at this level will *always* have enough space available in their reservation for this as they have already reserved the maximal amount of log metadata space they require, and such a delta reservation will always be less than or equal to the maximal amount in the reservation.

Hence we can grow the checkpoint transaction reservation dynamically as items are added to the CIL and avoid the need for reserving and regranting log space up front. This avoids deadlocks and removes a blocking point from the checkpoint flush code.

As mentioned early, transactions can't grow to more than half the size of the log. Hence as part of the reservation growing, we need to also check the size of the reservation against the maximum allowed transaction size. If we reach the maximum threshold, we need to push the CIL to the log. This is effectively a ”background flush” and is done on demand. This is identical to a CIL push triggered by a log force, only that there is no waiting for the checkpoint commit to complete. This background push is checked and executed by transaction commit code.

If the transaction subsystem goes idle while we still have items in the CIL, they will be flushed by the periodic log force issued by the xfssyncd. This log force will push the CIL to disk, and if the transaction subsystem stays idle, allow the idle log to be covered (effectively marked clean) in exactly the same manner that is done for the existing logging method. A discussion point is whether this log force needs to be done more frequently than the current rate which is once every 30s.

### Delayed Logging: Log Item Pinning

Currently log items are pinned during transaction commit while the items are still locked. This happens just after the items are formatted, though it could be done any time before the items are unlocked. The result of this mechanism is that items get pinned once for every transaction that is committed to the log buffers. Hence items that are relogged in the log buffers will have a pin count for every outstanding transaction they were dirtied in. When each of these transactions is completed, they will unpin the item once. As a result, the item only becomes unpinned when all the transactions complete and there are no pending transactions. Thus the pinning and unpinning of a log item is symmetric as there is a 1:1 relationship with transaction commit and log item completion.

For delayed logging, however, we have an asymmetric transaction commit to completion relationship. Every time an object is relogged in the CIL it goes through the commit process without a corresponding completion being registered. That is, we now have a many-to-one relationship between transaction commit and log item completion. The result of this is that pinning and unpinning of the log items becomes unbalanced if we retain the ”pin on transaction commit, unpin on transaction completion” model.

To keep pin/unpin symmetry, the algorithm needs to change to a ”pin on insertion into the CIL, unpin on checkpoint completion”. In other words, the pinning and unpinning becomes symmetric around a checkpoint context. We have to pin the object the first time it is inserted into the CIL - if it is already in the CIL during a transaction commit, then we do not pin it again. Because there can be multiple outstanding checkpoint contexts, we can still see elevated pin counts, but as each checkpoint completes the pin count will retain the correct value according to its context.

Just to make matters slightly more complex, this checkpoint level context for the pin count means that the pinning of an item must take place under the CIL commit/flush lock. If we pin the object outside this lock, we cannot guarantee which context the pin count is associated with. This is because of the fact pinning the item is dependent on whether the item is present in the current CIL or not. If we don't pin the CIL first before we check and pin the object, we have a race with CIL being flushed between the check and the pin (or not pinning, as the case may be). Hence we must hold the CIL flush/commit lock to guarantee that we pin the items correctly.

### Delayed Logging: Concurrent Scalability

A fundamental requirement for the CIL is that accesses through transaction commits must scale to many concurrent commits. The current transaction commit code does not break down even when there are transactions coming from 2048 processors at once. The current transaction code does not go any faster than if there was only one CPU using it, but it does not slow down either.

As a result, the delayed logging transaction commit code needs to be designed for concurrency from the ground up. It is obvious that there are serialisation points in the design - the three important ones are:

1. Locking out new transaction commits while flushing the CIL
2. Adding items to the CIL and updating item space accounting
3. Checkpoint commit ordering

Looking at the transaction commit and CIL flushing interactions, it is clear that we have a many-to-one interaction here. That is, the only restriction on the number of concurrent transactions that can be trying to commit at once is the amount of space available in the log for their reservations. The practical limit here is in the order of several hundred concurrent transactions for a 128MB log, which means that it is generally one per CPU in a machine.

The amount of time a transaction commit needs to hold out a flush is a relatively long period of time - the pinning of log items needs to be done while we are holding out a CIL flush, so at the moment that means it is held across the formatting of the objects into memory buffers (i.e. while memcpy()s are in progress). Ultimately a two pass algorithm where the formatting is done separately to the pinning of objects could be used to reduce the hold time of the transaction commit side.

Because of the number of potential transaction commit side holders, the lock really needs to be a sleeping lock - if the CIL flush takes the lock, we do not want every other CPU in the machine spinning on the CIL lock. Given that flushing the CIL could involve walking a list of tens of thousands of log items, it will get held for a significant time and so spin contention is a significant concern. Preventing lots of CPUs spinning doing nothing is the main reason for choosing a sleeping lock even though nothing in either the transaction commit or CIL flush side sleeps with the lock held.

It should also be noted that CIL flushing is also a relatively rare operation compared to transaction commit for asynchronous transaction workloads - only time will tell if using a read-write semaphore for exclusion will limit transaction commit concurrency due to cache line bouncing of the lock on the read side.

The second serialisation point is on the transaction commit side where items are inserted into the CIL. Because transactions can enter this code concurrently, the CIL needs to be protected separately from the above commit/flush exclusion. It also needs to be an exclusive lock but it is only held for a very short time and so a spin lock is appropriate here. It is possible that this lock will become a contention point, but given the short hold time once per transaction I think that contention is unlikely.

The final serialisation point is the checkpoint commit record ordering code that is run as part of the checkpoint commit and log force sequencing. The code path that triggers a CIL flush (i.e. whatever triggers the log force) will enter an ordering loop after writing all the log vectors into the log buffers but before writing the commit record. This loop walks the list of committing checkpoints and needs to block waiting for checkpoints to complete their commit record write. As a result it needs a lock and a wait variable. Log force sequencing also requires the same lock, list walk, and blocking mechanism to ensure completion of checkpoints.

These two sequencing operations can use the mechanism even though the events they are waiting for are different. The checkpoint commit record sequencing needs to wait until checkpoint contexts contain a commit LSN (obtained through completion of a commit record write) while log force sequencing needs to wait until previous checkpoint contexts are removed from the committing list (i.e. they've completed). A simple wait variable and broadcast wakeups (thundering herds) has been used to implement these two serialisation queues. They use the same lock as the CIL, too. If we see too much contention on the CIL lock, or too many context switches as a result of the broadcast wakeups these operations can be put under a new spinlock and given separate wait lists to reduce lock contention and the number of processes woken by the wrong event.

### Lifecycle Changes

The existing log item life cycle is as follows:

```
1. Transaction allocate
2. Transaction reserve
3. Lock item
4. Join item to transaction
        If not already attached,
                Allocate log item
                Attach log item to owner item
        Attach log item to transaction
5. Modify item
        Record modifications in log item
6. Transaction commit
        Pin item in memory
        Format item into log buffer
        Write commit LSN into transaction
        Unlock item
        Attach transaction to log buffer

<log buffer IO dispatched>
<log buffer IO completes>

7. Transaction completion
        Mark log item committed
        Insert log item into AIL
                Write commit LSN into log item
        Unpin log item
8. AIL traversal
        Lock item
        Mark log item clean
        Flush item to disk

<item IO completion>

9. Log item removed from AIL
        Moves log tail
        Item unlocked
```

Essentially, steps 1-6 operate independently from step 7, which is also independent of steps 8-9. An item can be locked in steps 1-6 or steps 8-9 at the same time step 7 is occurring, but only steps 1-6 or 8-9 can occur at the same time. If the log item is in the AIL or between steps 6 and 7 and steps 1-6 are re-entered, then the item is relogged. Only when steps 8-9 are entered and completed is the object considered clean.

With delayed logging, there are new steps inserted into the life cycle:

```
1. Transaction allocate
2. Transaction reserve
3. Lock item
4. Join item to transaction
```

```
        If not already attached,
                Allocate log item
                Attach log item to owner item
        Attach log item to transaction
5. Modify item
        Record modifications in log item
6. Transaction commit
        Pin item in memory if not pinned in CIL
        Format item into log vector + buffer
        Attach log vector and buffer to log item
        Insert log item into CIL
        Write CIL context sequence into transaction
        Unlock item

<next log force>

7. CIL push
        lock CIL flush
        Chain log vectors and buffers together
        Remove items from CIL
        unlock CIL flush
        write log vectors into log
        sequence commit records
        attach checkpoint context to log buffer

<log buffer IO dispatched>
<log buffer IO completes>

8. Checkpoint completion
        Mark log item committed
        Insert item into AIL
                Write commit LSN into log item
        Unpin log item
9. AIL traversal
        Lock item
        Mark log item clean
        Flush item to disk
<item IO completion>
10. Log item removed from AIL
        Moves log tail
        Item unlocked
```

From this, it can be seen that the only life cycle differences between the two logging methods are in the middle of the life cycle - they still have the same beginning and end and execution constraints. The only differences are in the committing of the log items to the log itself and the completion processing. Hence delayed logging should not introduce any constraints on log item behaviour, allocation or freeing that don't already exist.

As a result of this zero-impact ”insertion” of delayed logging infrastructure and the design of the internal structures to avoid on disk format changes, we can basically switch between delayed logging and the existing mechanism with a mount option. Fundamentally, there is no reason why

the log manager would not be able to swap methods automatically and transparently depending on load characteristics, but this should not be necessary if delayed logging works as designed.

### 3.59.2 XFS Maintainer Entry Profile

#### Overview

XFS is a well known high-performance filesystem in the Linux kernel. The aim of this project is to provide and maintain a robust and performant filesystem.

Patches are generally merged to the for-next branch of the appropriate git repository. After a testing period, the for-next branch is merged to the master branch.

Kernel code are merged to the xfs-linux tree[0]. Userspace code are merged to the xfsprogs tree[1]. Test cases are merged to the xfstests tree[2]. Ondisk format documentation are merged to the xfs-documentation tree[3].

All patchsets involving XFS *must* be cc'd in their entirety to the mailing list linux-xfs@vger.kernel.org.

#### Roles

There are eight key roles in the XFS project. A person can take on multiple roles, and a role can be filled by multiple people. Anyone taking on a role is advised to check in with themselves and others on a regular basis about burnout.

- **Outside Contributor**: Anyone who sends a patch but is not involved in the XFS project on a regular basis. These folks are usually people who work on other filesystems or elsewhere in the kernel community.
- **Developer**: Someone who is familiar with the XFS codebase enough to write new code, documentation, and tests.

  Developers can often be found in the IRC channel mentioned by the `C:` entry in the kernel MAINTAINERS file.
- **Senior Developer**: A developer who is very familiar with at least some part of the XFS codebase and/or other subsystems in the kernel. These people collectively decide the long term goals of the project and nudge the community in that direction. They should help prioritize development and review work for each release cycle.

  Senior developers tend to be more active participants in the IRC channel.
- **Reviewer**: Someone (most likely also a developer) who reads code submissions to decide:

  0. Is the idea behind the contribution sound?
  1. Does the idea fit the goals of the project?
  2. Is the contribution designed correctly?
  3. Is the contribution polished?
  4. Can the contribution be tested effectively?

  Reviewers should identify themselves with an `R:` entry in the kernel and fstests MAINTAINERS files.

- **Testing Lead**: This person is responsible for setting the test coverage goals of the project, negotiating with developers to decide on new tests for new features, and making sure that developers and release managers execute on the testing.

  The testing lead should identify themselves with an `M:` entry in the XFS section of the fstests MAINTAINERS file.

- **Bug Triager**: Someone who examines incoming bug reports in just enough detail to identify the person to whom the report should be forwarded.

  The bug triagers should identify themselves with a `B:` entry in the kernel MAINTAINERS file.

- **Release Manager**: This person merges reviewed patchsets into an integration branch, tests the result locally, pushes the branch to a public git repository, and sends pull requests further upstream. The release manager is not expected to work on new feature patchsets. If a developer and a reviewer fail to reach a resolution on some point, the release manager must have the ability to intervene to try to drive a resolution.

  The release manager should identify themselves with an `M:` entry in the kernel MAINTAINERS file.

- **Community Manager**: This person calls and moderates meetings of as many XFS participants as they can get when mailing list discussions prove insufficient for collective decisionmaking. They may also serve as liaison between managers of the organizations sponsoring work on any part of XFS.

- **LTS Maintainer**: Someone who backports and tests bug fixes from uptream to the LTS kernels. There tend to be six separate LTS trees at any given time.

  The maintainer for a given LTS release should identify themselves with an `M:` entry in the MAINTAINERS file for that LTS tree. Unmaintained LTS kernels should be marked with status `S: Orphan` in that same file.

### Submission Checklist Addendum

Please follow these additional rules when submitting to XFS:

- Patches affecting only the filesystem itself should be based against the latest -rc or the for-next branch. These patches will be merged back to the for-next branch.
- Authors of patches touching other subsystems need to coordinate with the maintainers of XFS and the relevant subsystems to decide how to proceed with a merge.
- Any patchset changing XFS should be cc'd in its entirety to linux-xfs. Do not send partial patchsets; that makes analysis of the broader context of the changes unnecessarily difficult.
- Anyone making kernel changes that have corresponding changes to the userspace utilities should send the userspace changes as separate patchsets immediately after the kernel patchsets.
- Authors of bug fix patches are expected to use fstests[2] to perform an A/B test of the patch to determine that there are no regressions. When possible, a new regression test case should be written for fstests.
- Authors of new feature patchsets must ensure that fstests will have appropriate functional and input corner-case test cases for the new feature.

- When implementing a new feature, it is strongly suggested that the developers write a design document to answer the following questions:
  - **What** problem is this trying to solve?
  - **Who** will benefit from this solution, and **where** will they access it?
  - **How** will this new feature work? This should touch on major data structures and algorithms supporting the solution at a higher level than code comments.
  - **What** userspace interfaces are necessary to build off of the new features?
  - **How** will this work be tested to ensure that it solves the problems laid out in the design document without causing new problems?

  The design document should be committed in the kernel documentation directory. It may be omitted if the feature is already well known to the community.
- Patchsets for the new tests should be submitted as separate patchsets immediately after the kernel and userspace code patchsets.
- Changes to the on-disk format of XFS must be described in the ondisk format document[3] and submitted as a patchset after the fstests patchsets.
- Patchsets implementing bug fixes and further code cleanups should put the bug fixes at the beginning of the series to ease backporting.

### Key Release Cycle Dates

Bug fixes may be sent at any time, though the release manager may decide to defer a patch when the next merge window is close.

Code submissions targeting the next merge window should be sent between -rc1 and -rc6. This gives the community time to review the changes, to suggest other changes, and for the author to retest those changes.

Code submissions also requiring changes to fs/iomap and targeting the next merge window should be sent between -rc1 and -rc4. This allows the broader kernel community adequate time to test the infrastructure changes.

### Review Cadence

In general, please wait at least one week before pinging for feedback. To find reviewers, either consult the MAINTAINERS file, or ask developers that have Reviewed-by tags for XFS changes to take a look and offer their opinion.

#### References

[0] https://git.kernel.org/pub/scm/fs/xfs/xfs-linux.git/
[1] https://git.kernel.org/pub/scm/fs/xfs/xfsprogs-dev.git/
[2] https://git.kernel.org/pub/scm/fs/xfs/xfstests-dev.git/
[3] https://git.kernel.org/pub/scm/fs/xfs/xfs-documentation.git/

### 3.59.3 XFS Self Describing Metadata

#### Introduction

The largest scalability problem facing XFS is not one of algorithmic scalability, but of verification of the filesystem structure. Scalabilty of the structures and indexes on disk and the algorithms for iterating them are adequate for supporting PB scale filesystems with billions of inodes, however it is this very scalability that causes the verification problem.

Almost all metadata on XFS is dynamically allocated. The only fixed location metadata is the allocation group headers (SB, AGF, AGFL and AGI), while all other metadata structures need to be discovered by walking the filesystem structure in different ways. While this is already done by userspace tools for validating and repairing the structure, there are limits to what they can verify, and this in turn limits the supportable size of an XFS filesystem.

For example, it is entirely possible to manually use xfs_db and a bit of scripting to analyse the structure of a 100TB filesystem when trying to determine the root cause of a corruption problem, but it is still mainly a manual task of verifying that things like single bit errors or misplaced writes weren't the ultimate cause of a corruption event. It may take a few hours to a few days to perform such forensic analysis, so for at this scale root cause analysis is entirely possible.

However, if we scale the filesystem up to 1PB, we now have 10x as much metadata to analyse and so that analysis blows out towards weeks/months of forensic work. Most of the analysis work is slow and tedious, so as the amount of analysis goes up, the more likely that the cause will be lost in the noise. Hence the primary concern for supporting PB scale filesystems is minimising the time and effort required for basic forensic analysis of the filesystem structure.

#### Self Describing Metadata

One of the problems with the current metadata format is that apart from the magic number in the metadata block, we have no other way of identifying what it is supposed to be. We can't even identify if it is the right place. Put simply, you can't look at a single metadata block in isolation and say ”yes, it is supposed to be there and the contents are valid”.

Hence most of the time spent on forensic analysis is spent doing basic verification of metadata values, looking for values that are in range (and hence not detected by automated verification checks) but are not correct. Finding and understanding how things like cross linked block lists (e.g. sibling pointers in a btree end up with loops in them) are the key to understanding what went wrong, but it is impossible to tell what order the blocks were linked into each other or written to disk after the fact.

Hence we need to record more information into the metadata to allow us to quickly determine if the metadata is intact and can be ignored for the purpose of analysis. We can't protect

against every possible type of error, but we can ensure that common types of errors are easily detectable. Hence the concept of self describing metadata.

The first, fundamental requirement of self describing metadata is that the metadata object contains some form of unique identifier in a well known location. This allows us to identify the expected contents of the block and hence parse and verify the metadata object. IF we can't independently identify the type of metadata in the object, then the metadata doesn't describe itself very well at all!

Luckily, almost all XFS metadata has magic numbers embedded already - only the AGFL, remote symlinks and remote attribute blocks do not contain identifying magic numbers. Hence we can change the on-disk format of all these objects to add more identifying information and detect this simply by changing the magic numbers in the metadata objects. That is, if it has the current magic number, the metadata isn't self identifying. If it contains a new magic number, it is self identifying and we can do much more expansive automated verification of the metadata object at runtime, during forensic analysis or repair.

As a primary concern, self describing metadata needs some form of overall integrity checking. We cannot trust the metadata if we cannot verify that it has not been changed as a result of external influences. Hence we need some form of integrity check, and this is done by adding CRC32c validation to the metadata block. If we can verify the block contains the metadata it was intended to contain, a large amount of the manual verification work can be skipped.

CRC32c was selected as metadata cannot be more than 64k in length in XFS and hence a 32 bit CRC is more than sufficient to detect multi-bit errors in metadata blocks. CRC32c is also now hardware accelerated on common CPUs so it is fast. So while CRC32c is not the strongest of possible integrity checks that could be used, it is more than sufficient for our needs and has relatively little overhead. Adding support for larger integrity fields and/or algorithms does really provide any extra value over CRC32c, but it does add a lot of complexity and so there is no provision for changing the integrity checking mechanism.

Self describing metadata needs to contain enough information so that the metadata block can be verified as being in the correct place without needing to look at any other metadata. This means it needs to contain location information. Just adding a block number to the metadata is not sufficient to protect against mis-directed writes - a write might be misdirected to the wrong LUN and so be written to the ”correct block” of the wrong filesystem. Hence location information must contain a filesystem identifier as well as a block number.

Another key information point in forensic analysis is knowing who the metadata block belongs to. We already know the type, the location, that it is valid and/or corrupted, and how long ago that it was last modified. Knowing the owner of the block is important as it allows us to find other related metadata to determine the scope of the corruption. For example, if we have a extent btree object, we don't know what inode it belongs to and hence have to walk the entire filesystem to find the owner of the block. Worse, the corruption could mean that no owner can be found (i.e. it's an orphan block), and so without an owner field in the metadata we have no idea of the scope of the corruption. If we have an owner field in the metadata object, we can immediately do top down validation to determine the scope of the problem.

Different types of metadata have different owner identifiers. For example, directory, attribute and extent tree blocks are all owned by an inode, while freespace btree blocks are owned by an allocation group. Hence the size and contents of the owner field are determined by the type of metadata object we are looking at. The owner information can also identify misplaced writes (e.g. freespace btree block written to the wrong AG).

Self describing metadata also needs to contain some indication of when it was written to the

filesystem. One of the key information points when doing forensic analysis is how recently the block was modified. Correlation of set of corrupted metadata blocks based on modification times is important as it can indicate whether the corruptions are related, whether there's been multiple corruption events that lead to the eventual failure, and even whether there are corruptions present that the run-time verification is not detecting.

For example, we can determine whether a metadata object is supposed to be free space or still allocated if it is still referenced by its owner by looking at when the free space btree block that contains the block was last written compared to when the metadata object itself was last written. If the free space block is more recent than the object and the object's owner, then there is a very good chance that the block should have been removed from the owner.

To provide this "written timestamp", each metadata block gets the Log Sequence Number (LSN) of the most recent transaction it was modified on written into it. This number will always increase over the life of the filesystem, and the only thing that resets it is running xfs_repair on the filesystem. Further, by use of the LSN we can tell if the corrupted metadata all belonged to the same log checkpoint and hence have some idea of how much modification occurred between the first and last instance of corrupt metadata on disk and, further, how much modification occurred between the corruption being written and when it was detected.

### Runtime Validation

Validation of self-describing metadata takes place at runtime in two places:

- immediately after a successful read from disk
- immediately prior to write IO submission

The verification is completely stateless - it is done independently of the modification process, and seeks only to check that the metadata is what it says it is and that the metadata fields are within bounds and internally consistent. As such, we cannot catch all types of corruption that can occur within a block as there may be certain limitations that operational state enforces of the metadata, or there may be corruption of interblock relationships (e.g. corrupted sibling pointer lists). Hence we still need stateful checking in the main code body, but in general most of the per-field validation is handled by the verifiers.

For read verification, the caller needs to specify the expected type of metadata that it should see, and the IO completion process verifies that the metadata object matches what was expected. If the verification process fails, then it marks the object being read as EFSCORRUPTED. The caller needs to catch this error (same as for IO errors), and if it needs to take special action due to a verification error it can do so by catching the EFSCORRUPTED error value. If we need more discrimination of error type at higher levels, we can define new error numbers for different errors as necessary.

The first step in read verification is checking the magic number and determining whether CRC validating is necessary. If it is, the CRC32c is calculated and compared against the value stored in the object itself. Once this is validated, further checks are made against the location information, followed by extensive object specific metadata validation. If any of these checks fail, then the buffer is considered corrupt and the EFSCORRUPTED error is set appropriately.

Write verification is the opposite of the read verification - first the object is extensively verified and if it is OK we then update the LSN from the last modification made to the object, After this, we calculate the CRC and insert it into the object. Once this is done the write IO is allowed to continue. If any error occurs during this process, the buffer is again marked with a EFSCORRUPTED error for the higher layers to catch.

### Structures

A typical on-disk structure needs to contain the following information:

```
struct xfs_ondisk_hdr {
        __be32  magic;              /* magic number */
        __be32  crc;                /* CRC, not logged */
        uuid_t  uuid;               /* filesystem identifier */
        __be64  owner;              /* parent object */
        __be64  blkno;              /* location on disk */
        __be64  lsn;                /* last modification in log, not logged */
};
```

Depending on the metadata, this information may be part of a header structure separate to the metadata contents, or may be distributed through an existing structure. The latter occurs with metadata that already contains some of this information, such as the superblock and AG headers.

Other metadata may have different formats for the information, but the same level of information is generally provided. For example:

- short btree blocks have a 32 bit owner (ag number) and a 32 bit block number for location. The two of these combined provide the same information as @owner and @blkno in eh above structure, but using 8 bytes less space on disk.
- directory/attribute node blocks have a 16 bit magic number, and the header that contains the magic number has other information in it as well. hence the additional metadata headers change the overall format of the metadata.

A typical buffer read verifier is structured as follows:

```
#define XFS_FOO_CRC_OFF             offsetof(struct xfs_ondisk_hdr, crc)

static void
xfs_foo_read_verify(
        struct xfs_buf      *bp)
{
    struct xfs_mount *mp = bp->b_mount;

        if ((xfs_sb_version_hascrc(&mp->m_sb) &&
            !xfs_verify_cksum(bp->b_addr, BBTOB(bp->b_length),
                                      XFS_FOO_CRC_OFF)) ||
            !xfs_foo_verify(bp)) {
                XFS_CORRUPTION_ERROR(__func__, XFS_ERRLEVEL_LOW, mp, bp->b_
↪addr);
                xfs_buf_ioerror(bp, EFSCORRUPTED);
        }
}
```

The code ensures that the CRC is only checked if the filesystem has CRCs enabled by checking the superblock of the feature bit, and then if the CRC verifies OK (or is not needed) it verifies the actual contents of the block.

The verifier function will take a couple of different forms, depending on whether the magic

number can be used to determine the format of the block. In the case it can't, the code is structured as follows:

```
static bool
xfs_foo_verify(
        struct xfs_buf              *bp)
{
        struct xfs_mount    *mp = bp->b_mount;
        struct xfs_ondisk_hdr       *hdr = bp->b_addr;

        if (hdr->magic != cpu_to_be32(XFS_FOO_MAGIC))
                return false;

        if (!xfs_sb_version_hascrc(&mp->m_sb)) {
                if (!uuid_equal(&hdr->uuid, &mp->m_sb.sb_uuid))
                        return false;
                if (bp->b_bn != be64_to_cpu(hdr->blkno))
                        return false;
                if (hdr->owner == 0)
                        return false;
        }

        /* object specific verification checks here */

        return true;
}
```

If there are different magic numbers for the different formats, the verifier will look like:

```
static bool
xfs_foo_verify(
        struct xfs_buf              *bp)
{
        struct xfs_mount    *mp = bp->b_mount;
        struct xfs_ondisk_hdr       *hdr = bp->b_addr;

        if (hdr->magic == cpu_to_be32(XFS_FOO_CRC_MAGIC)) {
                if (!uuid_equal(&hdr->uuid, &mp->m_sb.sb_uuid))
                        return false;
                if (bp->b_bn != be64_to_cpu(hdr->blkno))
                        return false;
                if (hdr->owner == 0)
                        return false;
        } else if (hdr->magic != cpu_to_be32(XFS_FOO_MAGIC))
                return false;

        /* object specific verification checks here */

        return true;
}
```

Write verifiers are very similar to the read verifiers, they just do things in the opposite order to

the read verifiers. A typical write verifier:

```
static void
xfs_foo_write_verify(
        struct xfs_buf      *bp)
{
        struct xfs_mount    *mp = bp->b_mount;
        struct xfs_buf_log_item     *bip = bp->b_fspriv;

        if (!xfs_foo_verify(bp)) {
                XFS_CORRUPTION_ERROR(__func__, XFS_ERRLEVEL_LOW, mp, bp->b_
↪addr);
                xfs_buf_ioerror(bp, EFSCORRUPTED);
                return;
        }

        if (!xfs_sb_version_hascrc(&mp->m_sb))
                return;


        if (bip) {
                struct xfs_ondisk_hdr       *hdr = bp->b_addr;
                hdr->lsn = cpu_to_be64(bip->bli_item.li_lsn);
        }
        xfs_update_cksum(bp->b_addr, BBTOB(bp->b_length), XFS_FOO_CRC_OFF);
}
```

This will verify the internal structure of the metadata before we go any further, detecting corruptions that have occurred as the metadata has been modified in memory. If the metadata verifies OK, and CRCs are enabled, we then update the LSN field (when it was last modified) and calculate the CRC on the metadata. Once this is done, we can issue the IO.

### Inodes and Dquots

Inodes and dquots are special snowflakes. They have per-object CRC and self-identifiers, but they are packed so that there are multiple objects per buffer. Hence we do not use per-buffer verifiers to do the work of per-object verification and CRC calculations. The per-buffer verifiers simply perform basic identification of the buffer - that they contain inodes or dquots, and that there are magic numbers in all the expected spots. All further CRC and verification checks are done when each inode is read from or written back to the buffer.

The structure of the verifiers and the identifiers checks is very similar to the buffer code described above. The only difference is where they are called. For example, inode read verification is done in xfs_inode_from_disk() when the inode is first read out of the buffer and the struct xfs_inode is instantiated. The inode is already extensively verified during writeback in xfs_iflush_int, so the only addition here is to add the LSN and CRC to the inode as it is copied back into the buffer.

XXX: inode unlinked list modification doesn't recalculate the inode CRC! None of the unlinked list modifications check or update CRCs, neither during unlink nor log recovery. So, it's gone unnoticed until now. This won't matter immediately - repair will probably complain about it - but it needs to be fixed.

### 3.59.4 XFS Online Fsck Design

This document captures the design of the online filesystem check feature for XFS. The purpose of this document is threefold:

- To help kernel distributors understand exactly what the XFS online fsck feature is, and issues about which they should be aware.
- To help people reading the code to familiarize themselves with the relevant concepts and design points before they start digging into the code.
- To help developers maintaining the system by capturing the reasons supporting higher level decision making.

As the online fsck code is merged, the links in this document to topic branches will be replaced with links to code.

This document is licensed under the terms of the GNU Public License, v2. The primary author is Darrick J. Wong.

This design document is split into seven parts. Part 1 defines what fsck tools are and the motivations for writing a new one. Parts 2 and 3 present a high level overview of how online fsck process works and how it is tested to ensure correct functionality. Part 4 discusses the user interface and the intended usage modes of the new program. Parts 5 and 6 show off the high level components and how they fit together, and then present case studies of how each repair function actually works. Part 7 sums up what has been discussed so far and speculates about what else might be built atop online fsck.

**Table of Contents**

## 1. What is a Filesystem Check?

A Unix filesystem has four main responsibilities:

- Provide a hierarchy of names through which application programs can associate arbitrary blobs of data for any length of time,
- Virtualize physical storage media across those names, and
- Retrieve the named data blobs at any time.
- Examine resource usage.

Metadata directly supporting these functions (e.g. files, directories, space mappings) are sometimes called primary metadata. Secondary metadata (e.g. reverse mapping and directory parent pointers) support operations internal to the filesystem, such as internal consistency check-

ing and reorganization. Summary metadata, as the name implies, condense information contained in primary metadata for performance reasons.

The filesystem check (fsck) tool examines all the metadata in a filesystem to look for errors. In addition to looking for obvious metadata corruptions, fsck also cross-references different types of metadata records with each other to look for inconsistencies. People do not like losing data, so most fsck tools also contains some ability to correct any problems found. As a word of caution -- the primary goal of most Linux fsck tools is to restore the filesystem metadata to a consistent state, not to maximize the data recovered. That precedent will not be challenged here.

Filesystems of the 20th century generally lacked any redundancy in the ondisk format, which means that fsck can only respond to errors by erasing files until errors are no longer detected. More recent filesystem designs contain enough redundancy in their metadata that it is now possible to regenerate data structures when non-catastrophic errors occur; this capability aids both strategies.

| **Note**: |
| --- |
| System administrators avoid data loss by increasing the number of separate storage systems through the creation of backups; and they avoid downtime by increasing the redundancy of each storage system through the creation of RAID arrays. fsck tools address only the first problem. |

### TLDR; Show Me the Code!

Code is posted to the kernel.org git trees as follows: kernel changes, userspace changes, and QA test changes. Each kernel patchset adding an online repair function will use the same branch name across the kernel, xfsprogs, and fstests git repos.

### Existing Tools

The online fsck tool described here will be the third tool in the history of XFS (on Linux) to check and repair filesystems. Two programs precede it:

The first program, `xfs_check`, was created as part of the XFS debugger (`xfs_db`) and can only be used with unmounted filesystems. It walks all metadata in the filesystem looking for inconsistencies in the metadata, though it lacks any ability to repair what it finds. Due to its high memory requirements and inability to repair things, this program is now deprecated and will not be discussed further.

The second program, `xfs_repair`, was created to be faster and more robust than the first program. Like its predecessor, it can only be used with unmounted filesystems. It uses extent-based in-memory data structures to reduce memory consumption, and tries to schedule readahead IO appropriately to reduce I/O waiting time while it scans the metadata of the entire filesystem. The most important feature of this tool is its ability to respond to inconsistencies in file metadata and directory tree by erasing things as needed to eliminate problems. Space usage metadata are rebuilt from the observed file metadata.

### Problem Statement

The current XFS tools leave several problems unsolved:

1. **User programs** suddenly **lose access** to the filesystem when unexpected shutdowns occur as a result of silent corruptions in the metadata. These occur **unpredictably** and often without warning.
2. **Users** experience a **total loss of service** during the recovery period after an **unexpected shutdown** occurs.
3. **Users** experience a **total loss of service** if the filesystem is taken offline to **look for problems** proactively.
4. **Data owners** cannot **check the integrity** of their stored data without reading all of it. This may expose them to substantial billing costs when a linear media scan performed by the storage system administrator might suffice.
5. **System administrators** cannot **schedule** a maintenance window to deal with corruptions if they **lack the means** to assess filesystem health while the filesystem is online.
6. **Fleet monitoring tools** cannot **automate periodic checks** of filesystem health when doing so requires **manual intervention** and downtime.
7. **Users** can be tricked into **doing things they do not desire** when malicious actors **exploit quirks of Unicode** to place misleading names in directories.

Given this definition of the problems to be solved and the actors who would benefit, the proposed solution is a third fsck tool that acts on a running filesystem.

This new third program has three components: an in-kernel facility to check metadata, an in-kernel facility to repair metadata, and a userspace driver program to drive fsck activity on a live filesystem. `xfs_scrub` is the name of the driver program. The rest of this document presents the goals and use cases of the new fsck tool, describes its major design points in connection to those goals, and discusses the similarities and differences with existing tools.

| **Note**: |
| --- |
| Throughout this document, the existing offline fsck tool can also be referred to by its current name "`xfs_repair`". The userspace driver program for the new online fsck tool can be referred to as "`xfs_scrub`". The kernel portion of online fsck that validates metadata is called "online scrub", and portion of the kernel that fixes metadata is called "online repair". |

The naming hierarchy is broken up into objects known as directories and files and the physical space is split into pieces known as allocation groups. Sharding enables better performance on highly parallel systems and helps to contain the damage when corruptions occur. The division of the filesystem into principal objects (allocation groups and inodes) means that there are ample opportunities to perform targeted checks and repairs on a subset of the filesystem.

While this is going on, other parts continue processing IO requests. Even if a piece of filesystem metadata can only be regenerated by scanning the entire system, the scan can still be done in the background while other file operations continue.

In summary, online fsck takes advantage of resource sharding and redundant metadata to enable targeted checking and repair operations while the system is running. This capability will be coupled to automatic system management so that autonomous self-healing of XFS maximizes service availability.

## 2. Theory of Operation

Because it is necessary for online fsck to lock and scan live metadata objects, online fsck consists of three separate code components. The first is the userspace driver program `xfs_scrub`, which is responsible for identifying individual metadata items, scheduling work items for them, reacting to the outcomes appropriately, and reporting results to the system administrator. The second and third are in the kernel, which implements functions to check and repair each type of online fsck work item.

| **Note**: |
| --- |
| For brevity, this document shortens the phrase ”online fsck work item” to ”scrub item”. |

Scrub item types are delineated in a manner consistent with the Unix design philosophy, which is to say that each item should handle one aspect of a metadata structure, and handle it well.

### Scope

In principle, online fsck should be able to check and to repair everything that the offline fsck program can handle. However, online fsck cannot be running 100% of the time, which means that latent errors may creep in after a scrub completes. If these errors cause the next mount to fail, offline fsck is the only solution. This limitation means that maintenance of the offline fsck tool will continue. A second limitation of online fsck is that it must follow the same resource sharing and lock acquisition rules as the regular filesystem. This means that scrub cannot take *any* shortcuts to save time, because doing so could lead to concurrency problems. In other words, online fsck is not a complete replacement for offline fsck, and a complete run of online fsck may take longer than online fsck. However, both of these limitations are acceptable tradeoffs to satisfy the different motivations of online fsck, which are to **minimize system downtime** and to **increase predictability of operation**.

### Phases of Work

The userspace driver program `xfs_scrub` splits the work of checking and repairing an entire filesystem into seven phases. Each phase concentrates on checking specific types of scrub items and depends on the success of all previous phases. The seven phases are as follows:

1. Collect geometry information about the mounted filesystem and computer, discover the online fsck capabilities of the kernel, and open the underlying storage devices.

2. Check allocation group metadata, all realtime volume metadata, and all quota files. Each metadata structure is scheduled as a separate scrub item. If corruption is found in the inode header or inode btree and `xfs_scrub` is permitted to perform repairs, then those scrub items are repaired to prepare for phase 3. Repairs are implemented by using the information in the scrub item to resubmit the kernel scrub call with the repair flag enabled; this is discussed in the next section. Optimizations and all other repairs are deferred to phase 4.

3. Check all metadata of every file in the filesystem. Each metadata structure is also scheduled as a separate scrub item. If repairs are needed and `xfs_scrub` is permitted to perform repairs, and there were no problems detected during phase 2, then those scrub items are repaired immediately. Optimizations, deferred repairs, and unsuccessful repairs are deferred to phase 4.

4. All remaining repairs and scheduled optimizations are performed during this phase, if the caller permits them. Before starting repairs, the summary counters are checked and any necessary repairs are performed so that subsequent repairs will not fail the resource reservation step due to wildly incorrect summary counters. Unsuccessful repairs are requeued as long as forward progress on repairs is made somewhere in the filesystem. Free space in the filesystem is trimmed at the end of phase 4 if the filesystem is clean.
5. By the start of this phase, all primary and secondary filesystem metadata must be correct. Summary counters such as the free space counts and quota resource counts are checked and corrected. Directory entry names and extended attribute names are checked for suspicious entries such as control characters or confusing Unicode sequences appearing in names.
6. If the caller asks for a media scan, read all allocated and written data file extents in the filesystem. The ability to use hardware-assisted data file integrity checking is new to online fsck; neither of the previous tools have this capability. If media errors occur, they will be mapped to the owning files and reported.
7. Re-check the summary counters and presents the caller with a summary of space usage and file counts.

This allocation of responsibilities will be *revisited* later in this document.

### Steps for Each Scrub Item

The kernel scrub code uses a three-step strategy for checking and repairing the one aspect of a metadata object represented by a scrub item:

1. The scrub item of interest is checked for corruptions; opportunities for optimization; and for values that are directly controlled by the system administrator but look suspicious. If the item is not corrupt or does not need optimization, resource are released and the positive scan results are returned to userspace. If the item is corrupt or could be optimized but the caller does not permit this, resources are released and the negative scan results are returned to userspace. Otherwise, the kernel moves on to the second step.
2. The repair function is called to rebuild the data structure. Repair functions generally choose rebuild a structure from other metadata rather than try to salvage the existing structure. If the repair fails, the scan results from the first step are returned to userspace. Otherwise, the kernel moves on to the third step.
3. In the third step, the kernel runs the same checks over the new metadata item to assess the efficacy of the repairs. The results of the reassessment are returned to userspace.

### Classification of Metadata

Each type of metadata object (and therefore each type of scrub item) is classified as follows:

### Primary Metadata

Metadata structures in this category should be most familiar to filesystem users either because they are directly created by the user or they index objects created by the user Most filesystem objects fall into this class:

- Free space and reference count information
- Inode records and indexes
- Storage mapping information for file data
- Directories
- Extended attributes
- Symbolic links
- Quota limits

Scrub obeys the same rules as regular filesystem accesses for resource and lock acquisition.

Primary metadata objects are the simplest for scrub to process. The principal filesystem object (either an allocation group or an inode) that owns the item being scrubbed is locked to guard against concurrent updates. The check function examines every record associated with the type for obvious errors and cross-references healthy records against other metadata to look for inconsistencies. Repairs for this class of scrub item are simple, since the repair function starts by holding all the resources acquired in the previous step. The repair function scans available metadata as needed to record all the observations needed to complete the structure. Next, it stages the observations in a new ondisk structure and commits it atomically to complete the repair. Finally, the storage from the old data structure are carefully reaped.

Because `xfs_scrub` locks a primary object for the duration of the repair, this is effectively an offline repair operation performed on a subset of the filesystem. This minimizes the complexity of the repair code because it is not necessary to handle concurrent updates from other threads, nor is it necessary to access any other part of the filesystem. As a result, indexed structures can be rebuilt very quickly, and programs trying to access the damaged structure will be blocked until repairs complete. The only infrastructure needed by the repair code are the staging area for observations and a means to write new structures to disk. Despite these limitations, the advantage that online repair holds is clear: targeted work on individual shards of the filesystem avoids total loss of service.

This mechanism is described in section 2.1 (”Off-Line Algorithm”) of V. Srinivasan and M. J. Carey, ”Performance of On-Line Index Construction Algorithms”, *Extending Database Technology*, pp. 293-309, 1992.

Most primary metadata repair functions stage their intermediate results in an in-memory array prior to formatting the new ondisk structure, which is very similar to the list-based algorithm discussed in section 2.3 (”List-Based Algorithms”) of Srinivasan. However, any data structure builder that maintains a resource lock for the duration of the repair is *always* an offline algorithm.

### Secondary Metadata

Metadata structures in this category reflect records found in primary metadata, but are only needed for online fsck or for reorganization of the filesystem.

Secondary metadata include:

- Reverse mapping information
- Directory parent pointers

This class of metadata is difficult for scrub to process because scrub attaches to the secondary object but needs to check primary metadata, which runs counter to the usual order of resource acquisition. Frequently, this means that full filesystems scans are necessary to rebuild the metadata. Check functions can be limited in scope to reduce runtime. Repairs, however, require a full scan of primary metadata, which can take a long time to complete. Under these conditions, `xfs_scrub` cannot lock resources for the entire duration of the repair.

Instead, repair functions set up an in-memory staging structure to store observations. Depending on the requirements of the specific repair function, the staging index will either have the same format as the ondisk structure or a design specific to that repair function. The next step is to release all locks and start the filesystem scan. When the repair scanner needs to record an observation, the staging data are locked long enough to apply the update. While the filesystem scan is in progress, the repair function hooks the filesystem so that it can apply pending filesystem updates to the staging information. Once the scan is done, the owning object is re-locked, the live data is used to write a new ondisk structure, and the repairs are committed atomically. The hooks are disabled and the staging staging area is freed. Finally, the storage from the old data structure are carefully reaped.

Introducing concurrency helps online repair avoid various locking problems, but comes at a high cost to code complexity. Live filesystem code has to be hooked so that the repair function can observe updates in progress. The staging area has to become a fully functional parallel structure so that updates can be merged from the hooks. Finally, the hook, the filesystem scan, and the inode locking model must be sufficiently well integrated that a hook event can decide if a given update should be applied to the staging structure.

In theory, the scrub implementation could apply these same techniques for primary metadata, but doing so would make it massively more complex and less performant. Programs attempting to access the damaged structures are not blocked from operation, which may cause application failure or an unplanned filesystem shutdown.

Inspiration for the secondary metadata repair strategy was drawn from section 2.4 of Srinivasan above, and sections 2 (”NSF: Inded Build Without Side-File”) and 3.1.1 (”Duplicate Key Insert Problem”) in C. Mohan, ”Algorithms for Creating Indexes for Very Large Tables Without Quiescing Updates”, 1992.

The sidecar index mentioned above bears some resemblance to the side file method mentioned in Srinivasan and Mohan. Their method consists of an index builder that extracts relevant record data to build the new structure as quickly as possible; and an auxiliary structure that captures all updates that would be committed to the index by other threads were the new index already online. After the index building scan finishes, the updates recorded in the side file are applied to the new index. To avoid conflicts between the index builder and other writer threads, the builder maintains a publicly visible cursor that tracks the progress of the scan through the record space. To avoid duplication of work between the side file and the index builder, side file updates are elided when the record ID for the update is greater than the cursor position within

the record ID space.

To minimize changes to the rest of the codebase, XFS online repair keeps the replacement index hidden until it's completely ready to go. In other words, there is no attempt to expose the keyspace of the new index while repair is running. The complexity of such an approach would be very high and perhaps more appropriate to building *new* indices.

**Future Work Question**: Can the full scan and live update code used to facilitate a repair also be used to implement a comprehensive check?

*Answer*: In theory, yes. Check would be much stronger if each scrub function employed these live scans to build a shadow copy of the metadata and then compared the shadow records to the ondisk records. However, doing that is a fair amount more work than what the checking functions do now. The live scans and hooks were developed much later. That in turn increases the runtime of those scrub functions.

### Summary Information

Metadata structures in this last category summarize the contents of primary metadata records. These are often used to speed up resource usage queries, and are many times smaller than the primary metadata which they represent.

Examples of summary information include:

- Summary counts of free space and inodes
- File link counts from directories
- Quota resource usage counts

Check and repair require full filesystem scans, but resource and lock acquisition follow the same paths as regular filesystem accesses.

The superblock summary counters have special requirements due to the underlying implementation of the incore counters, and will be treated separately. Check and repair of the other types of summary counters (quota resource counts and file link counts) employ the same filesystem scanning and hooking techniques as outlined above, but because the underlying data are sets of integer counters, the staging data need not be a fully functional mirror of the ondisk structure.

Inspiration for quota and file link count repair strategies were drawn from sections 2.12 (”Online Index Operations”) through 2.14 (”Incremental View Maintenance”) of G. Graefe, ”Concurrent Queries and Updates in Summary Views and Their Indexes”, 2011.

Since quotas are non-negative integer counts of resource usage, online quotacheck can use the incremental view deltas described in section 2.14 to track pending changes to the block and inode usage counts in each transaction, and commit those changes to a dquot side file when the transaction commits. Delta tracking is necessary for dquots because the index builder scans inodes, whereas the data structure being rebuilt is an index of dquots. Link count checking combines the view deltas and commit step into one because it sets attributes of the objects being scanned instead of writing them to a separate data structure. Each online fsck function will be discussed as case studies later in this document.

### Risk Management

During the development of online fsck, several risk factors were identified that may make the feature unsuitable for certain distributors and users. Steps can be taken to mitigate or eliminate those risks, though at a cost to functionality.

- **Decreased performance**: Adding metadata indices to the filesystem increases the time cost of persisting changes to disk, and the reverse space mapping and directory parent pointers are no exception. System administrators who require the maximum performance can disable the reverse mapping features at format time, though this choice dramatically reduces the ability of online fsck to find inconsistencies and repair them.
- **Incorrect repairs**: As with all software, there might be defects in the software that result in incorrect repairs being written to the filesystem. Systematic fuzz testing (detailed in the next section) is employed by the authors to find bugs early, but it might not catch everything. The kernel build system provides Kconfig options (`CONFIG_XFS_ONLINE_SCRUB` and `CONFIG_XFS_ONLINE_REPAIR`) to enable distributors to choose not to accept this risk. The xfsprogs build system has a configure option (`--enable-scrub=no`) that disables building of the `xfs_scrub` binary, though this is not a risk mitigation if the kernel functionality remains enabled.
- **Inability to repair**: Sometimes, a filesystem is too badly damaged to be repairable. If the keyspaces of several metadata indices overlap in some manner but a coherent narrative cannot be formed from records collected, then the repair fails. To reduce the chance that a repair will fail with a dirty transaction and render the filesystem unusable, the online repair functions have been designed to stage and validate all new records before committing the new structure.
- **Misbehavior**: Online fsck requires many privileges -- raw IO to block devices, opening files by handle, ignoring Unix discretionary access control, and the ability to perform administrative changes. Running this automatically in the background scares people, so the systemd background service is configured to run with only the privileges required. Obviously, this cannot address certain problems like the kernel crashing or deadlocking, but it should be sufficient to prevent the scrub process from escaping and reconfiguring the system. The cron job does not have this protection.
- **Fuzz Kiddiez**: There are many people now who seem to think that running automated fuzz testing of ondisk artifacts to find mischievous behavior and spraying exploit code onto the public mailing list for instant zero-day disclosure is somehow of some social benefit. In the view of this author, the benefit is realized only when the fuzz operators help to **fix** the flaws, but this opinion apparently is not widely shared among security ”researchers”. The XFS maintainers' continuing ability to manage these events presents an ongoing risk to the stability of the development process. Automated testing should front-load some of the risk while the feature is considered EXPERIMENTAL.

Many of these risks are inherent to software programming. Despite this, it is hoped that this new functionality will prove useful in reducing unexpected downtime.

## 3. Testing Plan

As stated before, fsck tools have three main goals:

1. Detect inconsistencies in the metadata;
2. Eliminate those inconsistencies; and
3. Minimize further loss of data.

Demonstrations of correct operation are necessary to build users' confidence that the software behaves within expectations. Unfortunately, it was not really feasible to perform regular exhaustive testing of every aspect of a fsck tool until the introduction of low-cost virtual machines with high-IOPS storage. With ample hardware availability in mind, the testing strategy for the online fsck project involves differential analysis against the existing fsck tools and systematic testing of every attribute of every type of metadata object. Testing can be split into four major categories, as discussed below.

### Integrated Testing with fstests

The primary goal of any free software QA effort is to make testing as inexpensive and widespread as possible to maximize the scaling advantages of community. In other words, testing should maximize the breadth of filesystem configuration scenarios and hardware setups. This improves code quality by enabling the authors of online fsck to find and fix bugs early, and helps developers of new features to find integration issues earlier in their development effort.

The Linux filesystem community shares a common QA testing suite, fstests, for functional and regression testing. Even before development work began on online fsck, fstests (when run on XFS) would run both the `xfs_check` and `xfs_repair -n` commands on the test and scratch filesystems between each test. This provides a level of assurance that the kernel and the fsck tools stay in alignment about what constitutes consistent metadata. During development of the online checking code, fstests was modified to run `xfs_scrub -n` between each test to ensure that the new checking code produces the same results as the two existing fsck tools.

To start development of online repair, fstests was modified to run `xfs_repair` to rebuild the filesystem's metadata indices between tests. This ensures that offline repair does not crash, leave a corrupt filesystem after it exists, or trigger complaints from the online check. This also established a baseline for what can and cannot be repaired offline. To complete the first phase of development of online repair, fstests was modified to be able to run `xfs_scrub` in a ”force rebuild” mode. This enables a comparison of the effectiveness of online repair as compared to the existing offline repair tools.

### General Fuzz Testing of Metadata Blocks

XFS benefits greatly from having a very robust debugging tool, `xfs_db`.

Before development of online fsck even began, a set of fstests were created to test the rather common fault that entire metadata blocks get corrupted. This required the creation of fstests library code that can create a filesystem containing every possible type of metadata object. Next, individual test cases were created to create a test filesystem, identify a single block of a specific type of metadata object, trash it with the existing `blocktrash` command in `xfs_db`, and test the reaction of a particular metadata validation strategy.

This earlier test suite enabled XFS developers to test the ability of the in-kernel validation functions and the ability of the offline fsck tool to detect and eliminate the inconsistent metadata. This part of the test suite was extended to cover online fsck in exactly the same manner.

In other words, for a given fstests filesystem configuration:

- For each metadata object existing on the filesystem:
  - Write garbage to it
  - Test the reactions of:
    1. The kernel verifiers to stop obviously bad metadata
    2. Offline repair (`xfs_repair`) to detect and fix
    3. Online repair (`xfs_scrub`) to detect and fix

### Targeted Fuzz Testing of Metadata Records

The testing plan for online fsck includes extending the existing fs testing infrastructure to provide a much more powerful facility: targeted fuzz testing of every metadata field of every metadata object in the filesystem. `xfs_db` can modify every field of every metadata structure in every block in the filesystem to simulate the effects of memory corruption and software bugs. Given that fstests already contains the ability to create a filesystem containing every metadata format known to the filesystem, `xfs_db` can be used to perform exhaustive fuzz testing!

For a given fstests filesystem configuration:

- For each metadata object existing on the filesystem...
  - For each record inside that metadata object...
    - For each field inside that record...
      - For each conceivable type of transformation that can be applied to a bit field...
        1. Clear all bits
        2. Set all bits
        3. Toggle the most significant bit
        4. Toggle the middle bit
        5. Toggle the least significant bit
        6. Add a small quantity
        7. Subtract a small quantity
        8. Randomize the contents
           - ...test the reactions of:
             1. The kernel verifiers to stop obviously bad metadata
             2. Offline checking (`xfs_repair -n`)
             3. Offline repair (`xfs_repair`)
             4. Online checking (`xfs_scrub -n`)

5. Online repair (`xfs_scrub`)
6. Both repair tools (`xfs_scrub` and then `xfs_repair` if online repair doesn't succeed)

This is quite the combinatoric explosion!

Fortunately, having this much test coverage makes it easy for XFS developers to check the responses of XFS' fsck tools. Since the introduction of the fuzz testing framework, these tests have been used to discover incorrect repair code and missing functionality for entire classes of metadata objects in `xfs_repair`. The enhanced testing was used to finalize the deprecation of `xfs_check` by confirming that `xfs_repair` could detect at least as many corruptions as the older tool.

These tests have been very valuable for `xfs_scrub` in the same ways -- they allow the online fsck developers to compare online fsck against offline fsck, and they enable XFS developers to find deficiencies in the code base.

Proposed patchsets include general fuzzer improvements, fuzzing baselines, and improvements in fuzz testing comprehensiveness.

### Stress Testing

A unique requirement to online fsck is the ability to operate on a filesystem concurrently with regular workloads. Although it is of course impossible to run `xfs_scrub` with *zero* observable impact on the running system, the online repair code should never introduce inconsistencies into the filesystem metadata, and regular workloads should never notice resource starvation. To verify that these conditions are being met, fstests has been enhanced in the following ways:

- For each scrub item type, create a test to exercise checking that item type while running `fsstress`.
- For each scrub item type, create a test to exercise repairing that item type while running `fsstress`.
- Race `fsstress` and `xfs_scrub -n` to ensure that checking the whole filesystem doesn't cause problems.
- Race `fsstress` and `xfs_scrub` in force-rebuild mode to ensure that force-repairing the whole filesystem doesn't cause problems.
- Race `xfs_scrub` in check and force-repair mode against `fsstress` while freezing and thawing the filesystem.
- Race `xfs_scrub` in check and force-repair mode against `fsstress` while remounting the filesystem read-only and read-write.
- The same, but running `fsx` instead of `fsstress`. (Not done yet?)

Success is defined by the ability to run all of these tests without observing any unexpected filesystem shutdowns due to corrupted metadata, kernel hang check warnings, or any other sort of mischief.

Proposed patchsets include general stress testing and the evolution of existing per-function stress testing.

## 4. User Interface

The primary user of online fsck is the system administrator, just like offline repair. Online fsck presents two modes of operation to administrators: A foreground CLI process for online fsck on demand, and a background service that performs autonomous checking and repair.

### Checking on Demand

For administrators who want the absolute freshest information about the metadata in a filesystem, `xfs_scrub` can be run as a foreground process on a command line. The program checks every piece of metadata in the filesystem while the administrator waits for the results to be reported, just like the existing `xfs_repair` tool. Both tools share a `-n` option to perform a read-only scan, and a `-v` option to increase the verbosity of the information reported.

A new feature of `xfs_scrub` is the `-x` option, which employs the error correction capabilities of the hardware to check data file contents. The media scan is not enabled by default because it may dramatically increase program runtime and consume a lot of bandwidth on older storage hardware.

The output of a foreground invocation is captured in the system log.

The `xfs_scrub_all` program walks the list of mounted filesystems and initiates `xfs_scrub` for each of them in parallel. It serializes scans for any filesystems that resolve to the same top level kernel block device to prevent resource overconsumption.

### Background Service

To reduce the workload of system administrators, the `xfs_scrub` package provides a suite of systemd timers and services that run online fsck automatically on weekends by default. The background service configures scrub to run with as little privilege as possible, the lowest CPU and IO priority, and in a CPU-constrained single threaded mode. This can be tuned by the systemd administrator at any time to suit the latency and throughput requirements of customer workloads.

The output of the background service is also captured in the system log. If desired, reports of failures (either due to inconsistencies or mere runtime errors) can be emailed automatically by setting the `EMAIL_ADDR` environment variable in the following service files:

- `xfs_scrub_fail@.service`
- `xfs_scrub_media_fail@.service`
- `xfs_scrub_all_fail.service`

The decision to enable the background scan is left to the system administrator. This can be done by enabling either of the following services:

- `xfs_scrub_all.timer` on systemd systems
- `xfs_scrub_all.cron` on non-systemd systems

This automatic weekly scan is configured out of the box to perform an additional media scan of all file data once per month. This is less foolproof than, say, storing file data block checksums, but much more performant if application software provides its own integrity checking,

redundancy can be provided elsewhere above the filesystem, or the storage device's integrity guarantees are deemed sufficient.

The systemd unit file definitions have been subjected to a security audit (as of systemd 249) to ensure that the xfs_scrub processes have as little access to the rest of the system as possible. This was performed via `systemd-analyze security`, after which privileges were restricted to the minimum required, sandboxing was set up to the maximal extent possible with sandboxing and system call filtering; and access to the filesystem tree was restricted to the minimum needed to start the program and access the filesystem being scanned. The service definition files restrict CPU usage to 80% of one CPU core, and apply as nice of a priority to IO and CPU scheduling as possible. This measure was taken to minimize delays in the rest of the filesystem. No such hardening has been performed for the cron job.

Proposed patchset: Enabling the xfs_scrub background service.

### Health Reporting

XFS caches a summary of each filesystem's health status in memory. The information is updated whenever `xfs_scrub` is run, or whenever inconsistencies are detected in the filesystem metadata during regular operations. System administrators should use the `health` command of `xfs_spaceman` to download this information into a human-readable format. If problems have been observed, the administrator can schedule a reduced service window to run the online repair tool to correct the problem. Failing that, the administrator can decide to schedule a maintenance window to run the traditional offline repair tool to correct the problem.

**Future Work Question**: Should the health reporting integrate with the new inotify fs error notification system? Would it be helpful for sysadmins to have a daemon to listen for corruption notifications and initiate a repair?

*Answer*: These questions remain unanswered, but should be a part of the conversation with early adopters and potential downstream users of XFS.

Proposed patchsets include wiring up health reports to correction returns and preservation of sickness info during memory reclaim.

## 5. Kernel Algorithms and Data Structures

This section discusses the key algorithms and data structures of the kernel code that provide the ability to check and repair metadata while the system is running. The first chapters in this section reveal the pieces that provide the foundation for checking metadata. The remainder of this section presents the mechanisms through which XFS regenerates itself.

### Self Describing Metadata

Starting with XFS version 5 in 2012, XFS updated the format of nearly every ondisk block header to record a magic number, a checksum, a universally "unique" identifier (UUID), an owner code, the ondisk address of the block, and a log sequence number. When loading a block buffer from disk, the magic number, UUID, owner, and ondisk address confirm that the retrieved block matches the specific owner of the current filesystem, and that the information contained in the block is supposed to be found at the ondisk address. The first three components enable

checking tools to disregard alleged metadata that doesn't belong to the filesystem, and the fourth component enables the filesystem to detect lost writes.

Whenever a file system operation modifies a block, the change is submitted to the log as part of a transaction. The log then processes these transactions marking them done once they are safely persisted to storage. The logging code maintains the checksum and the log sequence number of the last transactional update. Checksums are useful for detecting torn writes and other discrepancies that can be introduced between the computer and its storage devices. Sequence number tracking enables log recovery to avoid applying out of date log updates to the filesystem.

These two features improve overall runtime resiliency by providing a means for the filesystem to detect obvious corruption when reading metadata blocks from disk, but these buffer verifiers cannot provide any consistency checking between metadata structures.

For more information, please see the documentation for *XFS Self Describing Metadata*

### Reverse Mapping

The original design of XFS (circa 1993) is an improvement upon 1980s Unix filesystem design. In those days, storage density was expensive, CPU time was scarce, and excessive seek time could kill performance. For performance reasons, filesystem authors were reluctant to add redundancy to the filesystem, even at the cost of data integrity. Filesystems designers in the early 21st century choose different strategies to increase internal redundancy -- either storing nearly identical copies of metadata, or more space-efficient encoding techniques.

For XFS, a different redundancy strategy was chosen to modernize the design: a secondary space usage index that maps allocated disk extents back to their owners. By adding a new index, the filesystem retains most of its ability to scale well to heavily threaded workloads involving large datasets, since the primary file metadata (the directory tree, the file block map, and the allocation groups) remain unchanged. Like any system that improves redundancy, the reverse-mapping feature increases overhead costs for space mapping activities. However, it has two critical advantages: first, the reverse index is key to enabling online fsck and other requested functionality such as free space defragmentation, better media failure reporting, and filesystem shrinking. Second, the different ondisk storage format of the reverse mapping btree defeats device-level deduplication because the filesystem requires real redundancy.

| **Sidebar**: |
| --- |
| A criticism of adding the secondary index is that it does nothing to improve the robustness of user data storage itself. This is a valid point, but adding a new index for file data block checksums increases write amplification by turning data overwrites into copy-writes, which age the filesystem prematurely. In keeping with thirty years of precedent, users who want file data integrity can supply as powerful a solution as they require. As for metadata, the complexity of adding a new secondary index of space usage is much less than adding volume management and storage device mirroring to XFS itself. Perfection of RAID and volume management are best left to existing layers in the kernel. |

The information captured in a reverse space mapping record is as follows:

```
struct xfs_rmap_irec {
    xfs_agblock_t    rm_startblock;   /* extent start block */
    xfs_extlen_t     rm_blockcount;   /* extent length */
    uint64_t         rm_owner;        /* extent owner */
```

```
    uint64_t          rm_offset;      /* offset within the owner */
    unsigned int      rm_flags;       /* state flags */
};
```

The first two fields capture the location and size of the physical space, in units of filesystem blocks. The owner field tells scrub which metadata structure or file inode have been assigned this space. For space allocated to files, the offset field tells scrub where the space was mapped within the file fork. Finally, the flags field provides extra information about the space usage -- is this an attribute fork extent? A file mapping btree extent? Or an unwritten data extent?

Online filesystem checking judges the consistency of each primary metadata record by comparing its information against all other space indices. The reverse mapping index plays a key role in the consistency checking process because it contains a centralized alternate copy of all space allocation information. Program runtime and ease of resource acquisition are the only real limits to what online checking can consult. For example, a file data extent mapping can be checked against:

- The absence of an entry in the free space information.
- The absence of an entry in the inode index.
- The absence of an entry in the reference count data if the file is not marked as having shared extents.
- The correspondence of an entry in the reverse mapping information.

There are several observations to make about reverse mapping indices:

1. Reverse mappings can provide a positive affirmation of correctness if any of the above primary metadata are in doubt. The checking code for most primary metadata follows a path similar to the one outlined above.
2. Proving the consistency of secondary metadata with the primary metadata is difficult because that requires a full scan of all primary space metadata, which is very time intensive. For example, checking a reverse mapping record for a file extent mapping btree block requires locking the file and searching the entire btree to confirm the block. Instead, scrub relies on rigorous cross-referencing during the primary space mapping structure checks.
3. Consistency scans must use non-blocking lock acquisition primitives if the required locking order is not the same order used by regular filesystem operations. For example, if the filesystem normally takes a file ILOCK before taking the AGF buffer lock but scrub wants to take a file ILOCK while holding an AGF buffer lock, scrub cannot block on that second acquisition. This means that forward progress during this part of a scan of the reverse mapping data cannot be guaranteed if system load is heavy.

In summary, reverse mappings play a key role in reconstruction of primary metadata. The details of how these records are staged, written to disk, and committed into the filesystem are covered in subsequent sections.

### Checking and Cross-Referencing

The first step of checking a metadata structure is to examine every record contained within the structure and its relationship with the rest of the system. XFS contains multiple layers of checking to try to prevent inconsistent metadata from wreaking havoc on the system. Each of these layers contributes information that helps the kernel to make three decisions about the health of a metadata structure:

- Is a part of this structure obviously corrupt (`XFS_SCRUB_OFLAG_CORRUPT`) ?
- Is this structure inconsistent with the rest of the system (`XFS_SCRUB_OFLAG_XCORRUPT`) ?
- Is there so much damage around the filesystem that cross-referencing is not possible (`XFS_SCRUB_OFLAG_XFAIL`) ?
- Can the structure be optimized to improve performance or reduce the size of metadata (`XFS_SCRUB_OFLAG_PREEN`) ?
- Does the structure contain data that is not inconsistent but deserves review by the system administrator (`XFS_SCRUB_OFLAG_WARNING`) ?

The following sections describe how the metadata scrubbing process works.

### Metadata Buffer Verification

The lowest layer of metadata protection in XFS are the metadata verifiers built into the buffer cache. These functions perform inexpensive internal consistency checking of the block itself, and answer these questions:

- Does the block belong to this filesystem?
- Does the block belong to the structure that asked for the read? This assumes that metadata blocks only have one owner, which is always true in XFS.
- Is the type of data stored in the block within a reasonable range of what scrub is expecting?
- Does the physical location of the block match the location it was read from?
- Does the block checksum match the data?

The scope of the protections here are very limited -- verifiers can only establish that the filesystem code is reasonably free of gross corruption bugs and that the storage system is reasonably competent at retrieval. Corruption problems observed at runtime cause the generation of health reports, failed system calls, and in the extreme case, filesystem shutdowns if the corrupt metadata force the cancellation of a dirty transaction.

Every online fsck scrubbing function is expected to read every ondisk metadata block of a structure in the course of checking the structure. Corruption problems observed during a check are immediately reported to userspace as corruption; during a cross-reference, they are reported as a failure to cross-reference once the full examination is complete. Reads satisfied by a buffer already in cache (and hence already verified) bypass these checks.

### Internal Consistency Checks

After the buffer cache, the next level of metadata protection is the internal record verification code built into the filesystem. These checks are split between the buffer verifiers, the in-filesystem users of the buffer cache, and the scrub code itself, depending on the amount of higher level context required. The scope of checking is still internal to the block. These higher level checking functions answer these questions:

- Does the type of data stored in the block match what scrub is expecting?
- Does the block belong to the owning structure that asked for the read?
- If the block contains records, do the records fit within the block?
- If the block tracks internal free space information, is it consistent with the record areas?
- Are the records contained inside the block free of obvious corruptions?

Record checks in this category are more rigorous and more time-intensive. For example, block pointers and inumbers are checked to ensure that they point within the dynamically allocated parts of an allocation group and within the filesystem. Names are checked for invalid characters, and flags are checked for invalid combinations. Other record attributes are checked for sensible values. Btree records spanning an interval of the btree keyspace are checked for correct order and lack of mergeability (except for file fork mappings). For performance reasons, regular code may skip some of these checks unless debugging is enabled or a write is about to occur. Scrub functions, of course, must check all possible problems.

### Validation of Userspace-Controlled Record Attributes

Various pieces of filesystem metadata are directly controlled by userspace. Because of this nature, validation work cannot be more precise than checking that a value is within the possible range. These fields include:

- Superblock fields controlled by mount options
- Filesystem labels
- File timestamps
- File permissions
- File size
- File flags
- Names present in directory entries, extended attribute keys, and filesystem labels
- Extended attribute key namespaces
- Extended attribute values
- File data block contents
- Quota limits
- Quota timer expiration (if resource usage exceeds the soft limit)

### Cross-Referencing Space Metadata

After internal block checks, the next higher level of checking is cross-referencing records between metadata structures. For regular runtime code, the cost of these checks is considered to be prohibitively expensive, but as scrub is dedicated to rooting out inconsistencies, it must pursue all avenues of inquiry. The exact set of cross-referencing is highly dependent on the context of the data structure being checked.

The XFS btree code has keyspace scanning functions that online fsck uses to cross reference one structure with another. Specifically, scrub can scan the key space of an index to determine if that keyspace is fully, sparsely, or not at all mapped to records. For the reverse mapping btree, it is possible to mask parts of the key for the purposes of performing a keyspace scan so that scrub can decide if the rmap btree contains records mapping a certain extent of physical space without the sparsenses of the rest of the rmap keyspace getting in the way.

Btree blocks undergo the following checks before cross-referencing:

- Does the type of data stored in the block match what scrub is expecting?
- Does the block belong to the owning structure that asked for the read?
- Do the records fit within the block?
- Are the records contained inside the block free of obvious corruptions?
- Are the name hashes in the correct order?
- Do node pointers within the btree point to valid block addresses for the type of btree?
- Do child pointers point towards the leaves?
- Do sibling pointers point across the same level?
- For each node block record, does the record key accurate reflect the contents of the child block?

Space allocation records are cross-referenced as follows:

1. Any space mentioned by any metadata structure are cross-referenced as follows:
    - Does the reverse mapping index list only the appropriate owner as the owner of each block?
    - Are none of the blocks claimed as free space?
    - If these aren't file data blocks, are none of the blocks claimed as space shared by different owners?
2. Btree blocks are cross-referenced as follows:
    - Everything in class 1 above.
    - If there's a parent node block, do the keys listed for this block match the keyspace of this block?
    - Do the sibling pointers point to valid blocks? Of the same level?
    - Do the child pointers point to valid blocks? Of the next level down?
3. Free space btree records are cross-referenced as follows:
    - Everything in class 1 and 2 above.

- Does the reverse mapping index list no owners of this space?
- Is this space not claimed by the inode index for inodes?
- Is it not mentioned by the reference count index?
- Is there a matching record in the other free space btree?

4. Inode btree records are cross-referenced as follows:
   - Everything in class 1 and 2 above.
   - Is there a matching record in free inode btree?
   - Do cleared bits in the holemask correspond with inode clusters?
   - Do set bits in the freemask correspond with inode records with zero link count?
5. Inode records are cross-referenced as follows:
   - Everything in class 1.
   - Do all the fields that summarize information about the file forks actually match those forks?
   - Does each inode with zero link count correspond to a record in the free inode btree?
6. File fork space mapping records are cross-referenced as follows:
   - Everything in class 1 and 2 above.
   - Is this space not mentioned by the inode btrees?
   - If this is a CoW fork mapping, does it correspond to a CoW entry in the reference count btree?
7. Reference count records are cross-referenced as follows:
   - Everything in class 1 and 2 above.
   - Within the space subkeyspace of the rmap btree (that is to say, all records mapped to a particular space extent and ignoring the owner info), are there the same number of reverse mapping records for each block as the reference count record claims?

Proposed patchsets are the series to find gaps in refcount btree, inode btree, and rmap btree records; to find mergeable records; and to improve cross referencing with rmap before starting a repair.

### Checking Extended Attributes

Extended attributes implement a key-value store that enable fragments of data to be attached to any file. Both the kernel and userspace can access the keys and values, subject to namespace and privilege restrictions. Most typically these fragments are metadata about the file -- origins, security contexts, user-supplied labels, indexing information, etc.

Names can be as long as 255 bytes and can exist in several different namespaces. Values can be as large as 64KB. A file's extended attributes are stored in blocks mapped by the attr fork. The mappings point to leaf blocks, remote value blocks, or dabtree blocks. Block 0 in the attribute fork is always the top of the structure, but otherwise each of the three types of blocks can be found at any offset in the attr fork. Leaf blocks contain attribute key records that point to the name and the value. Names are always stored elsewhere in the same leaf block. Values that

are less than 3/4 the size of a filesystem block are also stored elsewhere in the same leaf block. Remote value blocks contain values that are too large to fit inside a leaf. If the leaf information exceeds a single filesystem block, a dabtree (also rooted at block 0) is created to map hashes of the attribute names to leaf blocks in the attr fork.

Checking an extended attribute structure is not so straightforward due to the lack of separation between attr blocks and index blocks. Scrub must read each block mapped by the attr fork and ignore the non-leaf blocks:

1. Walk the dabtree in the attr fork (if present) to ensure that there are no irregularities in the blocks or dabtree mappings that do not point to attr leaf blocks.
2. Walk the blocks of the attr fork looking for leaf blocks. For each entry inside a leaf:
    a. Validate that the name does not contain invalid characters.
    b. Read the attr value. This performs a named lookup of the attr name to ensure the correctness of the dabtree. If the value is stored in a remote block, this also validates the integrity of the remote value block.

### Checking and Cross-Referencing Directories

The filesystem directory tree is a directed acylic graph structure, with files constituting the nodes, and directory entries (dirents) constituting the edges. Directories are a special type of file containing a set of mappings from a 255-byte sequence (name) to an inumber. These are called directory entries, or dirents for short. Each directory file must have exactly one directory pointing to the file. A root directory points to itself. Directory entries point to files of any type. Each non-directory file may have multiple directories point to it.

In XFS, directories are implemented as a file containing up to three 32GB partitions. The first partition contains directory entry data blocks. Each data block contains variable-sized records associating a user-provided name with an inumber and, optionally, a file type. If the directory entry data grows beyond one block, the second partition (which exists as post-EOF extents) is populated with a block containing free space information and an index that maps hashes of the dirent names to directory data blocks in the first partition. This makes directory name lookups very fast. If this second partition grows beyond one block, the third partition is populated with a linear array of free space information for faster expansions. If the free space has been separated and the second partition grows again beyond one block, then a dabtree is used to map hashes of dirent names to directory data blocks.

Checking a directory is pretty straightforward:

1. Walk the dabtree in the second partition (if present) to ensure that there are no irregularities in the blocks or dabtree mappings that do not point to dirent blocks.
2. Walk the blocks of the first partition looking for directory entries. Each dirent is checked as follows:
    a. Does the name contain no invalid characters?
    b. Does the inumber correspond to an actual, allocated inode?
    c. Does the child inode have a nonzero link count?
    d. If a file type is included in the dirent, does it match the type of the inode?
    e. If the child is a subdirectory, does the child's dotdot pointer point back to the parent?

f. If the directory has a second partition, perform a named lookup of the dirent name to ensure the correctness of the dabtree.

3. Walk the free space list in the third partition (if present) to ensure that the free spaces it describes are really unused.

Checking operations involving *parents* and *file link counts* are discussed in more detail in later sections.

### Checking Directory/Attribute Btrees

As stated in previous sections, the directory/attribute btree (dabtree) index maps user-provided names to improve lookup times by avoiding linear scans. Internally, it maps a 32-bit hash of the name to a block offset within the appropriate file fork.

The internal structure of a dabtree closely resembles the btrees that record fixed-size metadata records -- each dabtree block contains a magic number, a checksum, sibling pointers, a UUID, a tree level, and a log sequence number. The format of leaf and node records are the same -- each entry points to the next level down in the hierarchy, with dabtree node records pointing to dabtree leaf blocks, and dabtree leaf records pointing to non-dabtree blocks elsewhere in the fork.

Checking and cross-referencing the dabtree is very similar to what is done for space btrees:

- Does the type of data stored in the block match what scrub is expecting?
- Does the block belong to the owning structure that asked for the read?
- Do the records fit within the block?
- Are the records contained inside the block free of obvious corruptions?
- Are the name hashes in the correct order?
- Do node pointers within the dabtree point to valid fork offsets for dabtree blocks?
- Do leaf pointers within the dabtree point to valid fork offsets for directory or attr leaf blocks?
- Do child pointers point towards the leaves?
- Do sibling pointers point across the same level?
- For each dabtree node record, does the record key accurate reflect the contents of the child dabtree block?
- For each dabtree leaf record, does the record key accurate reflect the contents of the directory or attr block?

### Cross-Referencing Summary Counters

XFS maintains three classes of summary counters: available resources, quota resource usage, and file link counts.

In theory, the amount of available resources (data blocks, inodes, realtime extents) can be found by walking the entire filesystem. This would make for very slow reporting, so a transactional filesystem can maintain summaries of this information in the superblock. Cross-referencing these values against the filesystem metadata should be a simple matter of walking the free space and inode metadata in each AG and the realtime bitmap, but there are complications that will be discussed in *more detail* later.

*Quota usage* and *file link count* checking are sufficiently complicated to warrant separate sections.

### Post-Repair Reverification

After performing a repair, the checking code is run a second time to validate the new structure, and the results of the health assessment are recorded internally and returned to the calling process. This step is critical for enabling system administrator to monitor the status of the filesystem and the progress of any repairs. For developers, it is a useful means to judge the efficacy of error detection and correction in the online and offline checking tools.

### Eventual Consistency vs. Online Fsck

Complex operations can make modifications to multiple per-AG data structures with a chain of transactions. These chains, once committed to the log, are restarted during log recovery if the system crashes while processing the chain. Because the AG header buffers are unlocked between transactions within a chain, online checking must coordinate with chained operations that are in progress to avoid incorrectly detecting inconsistencies due to pending chains. Furthermore, online repair must not run when operations are pending because the metadata are temporarily inconsistent with each other, and rebuilding is not possible.

Only online fsck has this requirement of total consistency of AG metadata, and should be relatively rare as compared to filesystem change operations. Online fsck coordinates with transaction chains as follows:

- For each AG, maintain a count of intent items targeting that AG. The count should be bumped whenever a new item is added to the chain. The count should be dropped when the filesystem has locked the AG header buffers and finished the work.
- When online fsck wants to examine an AG, it should lock the AG header buffers to quiesce all transaction chains that want to modify that AG. If the count is zero, proceed with the checking operation. If it is nonzero, cycle the buffer locks to allow the chain to make forward progress.

This may lead to online fsck taking a long time to complete, but regular filesystem updates take precedence over background checking activity. Details about the discovery of this situation are presented in the *next section*, and details about the solution are presented *after that*.

### Discovery of the Problem

Midway through the development of online scrubbing, the fsstress tests uncovered a misinteraction between online fsck and compound transaction chains created by other writer threads that resulted in false reports of metadata inconsistency. The root cause of these reports is the eventual consistency model introduced by the expansion of deferred work items and compound transaction chains when reverse mapping and reflink were introduced.

Originally, transaction chains were added to XFS to avoid deadlocks when unmapping space from files. Deadlock avoidance rules require that AGs only be locked in increasing order, which makes it impossible (say) to use a single transaction to free a space extent in AG 7 and then try to free a now superfluous block mapping btree block in AG 3. To avoid these kinds of deadlocks, XFS creates Extent Freeing Intent (EFI) log items to commit to freeing some space in one transaction while deferring the actual metadata updates to a fresh transaction. The transaction sequence looks like this:

1. The first transaction contains a physical update to the file's block mapping structures to remove the mapping from the btree blocks. It then attaches to the in-memory transaction an action item to schedule deferred freeing of space. Concretely, each transaction maintains a list of `struct xfs_defer_pending` objects, each of which maintains a list of `struct xfs_extent_free_item` objects. Returning to the example above, the action item tracks the freeing of both the unmapped space from AG 7 and the block mapping btree (BMBT) block from AG 3. Deferred frees recorded in this manner are committed in the log by creating an EFI log item from the `struct xfs_extent_free_item` object and attaching the log item to the transaction. When the log is persisted to disk, the EFI item is written into the ondisk transaction record. EFIs can list up to 16 extents to free, all sorted in AG order.
2. The second transaction contains a physical update to the free space btrees of AG 3 to release the former BMBT block and a second physical update to the free space btrees of AG 7 to release the unmapped file space. Observe that the physical updates are resequenced in the correct order when possible. Attached to the transaction is a an extent free done (EFD) log item. The EFD contains a pointer to the EFI logged in transaction #1 so that log recovery can tell if the EFI needs to be replayed.

If the system goes down after transaction #1 is written back to the filesystem but before #2 is committed, a scan of the filesystem metadata would show inconsistent filesystem metadata because there would not appear to be any owner of the unmapped space. Happily, log recovery corrects this inconsistency for us -- when recovery finds an intent log item but does not find a corresponding intent done item, it will reconstruct the incore state of the intent item and finish it. In the example above, the log must replay both frees described in the recovered EFI to complete the recovery phase.

There are subtleties to XFS' transaction chaining strategy to consider:

- Log items must be added to a transaction in the correct order to prevent conflicts with principal objects that are not held by the transaction. In other words, all per-AG metadata updates for an unmapped block must be completed before the last update to free the extent, and extents should not be reallocated until that last update commits to the log.
- AG header buffers are released between each transaction in a chain. This means that other threads can observe an AG in an intermediate state, but as long as the first subtlety is handled, this should not affect the correctness of filesystem operations.
- Unmounting the filesystem flushes all pending work to disk, which means that offline fsck

never sees the temporary inconsistencies caused by deferred work item processing.

In this manner, XFS employs a form of eventual consistency to avoid deadlocks and increase parallelism.

During the design phase of the reverse mapping and reflink features, it was decided that it was impractical to cram all the reverse mapping updates for a single filesystem change into a single transaction because a single file mapping operation can explode into many small updates:

- The block mapping update itself
- A reverse mapping update for the block mapping update
- Fixing the freelist
- A reverse mapping update for the freelist fix
- A shape change to the block mapping btree
- A reverse mapping update for the btree update
- Fixing the freelist (again)
- A reverse mapping update for the freelist fix
- An update to the reference counting information
- A reverse mapping update for the refcount update
- Fixing the freelist (a third time)
- A reverse mapping update for the freelist fix
- Freeing any space that was unmapped and not owned by any other file
- Fixing the freelist (a fourth time)
- A reverse mapping update for the freelist fix
- Freeing the space used by the block mapping btree
- Fixing the freelist (a fifth time)
- A reverse mapping update for the freelist fix

Free list fixups are not usually needed more than once per AG per transaction chain, but it is theoretically possible if space is very tight. For copy-on-write updates this is even worse, because this must be done once to remove the space from a staging area and again to map it into the file!

To deal with this explosion in a calm manner, XFS expands its use of deferred work items to cover most reverse mapping updates and all refcount updates. This reduces the worst case size of transaction reservations by breaking the work into a long chain of small updates, which increases the degree of eventual consistency in the system. Again, this generally isn't a problem because XFS orders its deferred work items carefully to avoid resource reuse conflicts between unsuspecting threads.

However, online fsck changes the rules -- remember that although physical updates to per-AG structures are coordinated by locking the buffers for AG headers, buffer locks are dropped between transactions. Once scrub acquires resources and takes locks for a data structure, it must do all the validation work without releasing the lock. If the main lock for a space btree is an AG header buffer lock, scrub may have interrupted another thread that is midway through finishing a chain. For example, if a thread performing a copy-on-write has completed a reverse

mapping update but not the corresponding refcount update, the two AG btrees will appear inconsistent to scrub and an observation of corruption will be recorded. This observation will not be correct. If a repair is attempted in this state, the results will be catastrophic!

Several other solutions to this problem were evaluated upon discovery of this flaw and rejected:

1. Add a higher level lock to allocation groups and require writer threads to acquire the higher level lock in AG order before making any changes. This would be very difficult to implement in practice because it is difficult to determine which locks need to be obtained, and in what order, without simulating the entire operation. Performing a dry run of a file operation to discover necessary locks would make the filesystem very slow.
2. Make the deferred work coordinator code aware of consecutive intent items targeting the same AG and have it hold the AG header buffers locked across the transaction roll between updates. This would introduce a lot of complexity into the coordinator since it is only loosely coupled with the actual deferred work items. It would also fail to solve the problem because deferred work items can generate new deferred subtasks, but all subtasks must be complete before work can start on a new sibling task.
3. Teach online fsck to walk all transactions waiting for whichever lock(s) protect the data structure being scrubbed to look for pending operations. The checking and repair operations must factor these pending operations into the evaluations being performed. This solution is a nonstarter because it is *extremely* invasive to the main filesystem.

### Intent Drains

Online fsck uses an atomic intent item counter and lock cycling to coordinate with transaction chains. There are two key properties to the drain mechanism. First, the counter is incremented when a deferred work item is *queued* to a transaction, and it is decremented after the associated intent done log item is *committed* to another transaction. The second property is that deferred work can be added to a transaction without holding an AG header lock, but per-AG work items cannot be marked done without locking that AG header buffer to log the physical updates and the intent done log item. The first property enables scrub to yield to running transaction chains, which is an explicit deprioritization of online fsck to benefit file operations. The second property of the drain is key to the correct coordination of scrub, since scrub will always be able to decide if a conflict is possible.

For regular filesystem code, the drain works as follows:

1. Call the appropriate subsystem function to add a deferred work item to a transaction.
2. The function calls `xfs_defer_drain_bump` to increase the counter.
3. When the deferred item manager wants to finish the deferred work item, it calls `->finish_item` to complete it.
4. The `->finish_item` implementation logs some changes and calls `xfs_defer_drain_drop` to decrease the sloppy counter and wake up any threads waiting on the drain.
5. The subtransaction commits, which unlocks the resource associated with the intent item.

For scrub, the drain works as follows:

1. Lock the resource(s) associated with the metadata being scrubbed. For example, a scan of the refcount btree would lock the AGI and AGF header buffers.

2. If the counter is zero (`xfs_defer_drain_busy` returns false), there are no chains in progress and the operation may proceed.
3. Otherwise, release the resources grabbed in step 1.
4. Wait for the intent counter to reach zero (`xfs_defer_drain_intents`), then go back to step 1 unless a signal has been caught.

To avoid polling in step 4, the drain provides a waitqueue for scrub threads to be woken up whenever the intent count drops to zero.

The proposed patchset is the scrub intent drain series.

### Static Keys (aka Jump Label Patching)

Online fsck for XFS separates the regular filesystem from the checking and repair code as much as possible. However, there are a few parts of online fsck (such as the intent drains, and later, live update hooks) where it is useful for the online fsck code to know what's going on in the rest of the filesystem. Since it is not expected that online fsck will be constantly running in the background, it is very important to minimize the runtime overhead imposed by these hooks when online fsck is compiled into the kernel but not actively running on behalf of userspace. Taking locks in the hot path of a writer thread to access a data structure only to find that no further action is necessary is expensive -- on the author's computer, this have an overhead of 40-50ns per access. Fortunately, the kernel supports dynamic code patching, which enables XFS to replace a static branch to hook code with `nop` sleds when online fsck isn't running. This sled has an overhead of however long it takes the instruction decoder to skip past the sled, which seems to be on the order of less than 1ns and does not access memory outside of instruction fetching.

When online fsck enables the static key, the sled is replaced with an unconditional branch to call the hook code. The switchover is quite expensive (~22000ns) but is paid entirely by the program that invoked online fsck, and can be amortized if multiple threads enter online fsck at the same time, or if multiple filesystems are being checked at the same time. Changing the branch direction requires taking the CPU hotplug lock, and since CPU initialization requires memory allocation, online fsck must be careful not to change a static key while holding any locks or resources that could be accessed in the memory reclaim paths. To minimize contention on the CPU hotplug lock, care should be taken not to enable or disable static keys unnecessarily.

Because static keys are intended to minimize hook overhead for regular filesystem operations when xfs_scrub is not running, the intended usage patterns are as follows:

- The hooked part of XFS should declare a static-scoped static key that defaults to false. The `DEFINE_STATIC_KEY_FALSE` macro takes care of this. The static key itself should be declared as a `static` variable.
- When deciding to invoke code that's only used by scrub, the regular filesystem should call the `static_branch_unlikely` predicate to avoid the scrub-only hook code if the static key is not enabled.
- The regular filesystem should export helper functions that call `static_branch_inc` to enable and `static_branch_dec` to disable the static key. Wrapper functions make it easy to compile out the relevant code if the kernel distributor turns off online fsck at build time.
- Scrub functions wanting to turn on scrub-only XFS functionality should call the `xchk_fsgates_enable` from the setup function to enable a specific hook. This must be

done before obtaining any resources that are used by memory reclaim. Callers had better be sure they really need the functionality gated by the static key; the `TRY_HARDER` flag is useful here.

Online scrub has resource acquisition helpers (e.g. `xchk_perag_lock`) to handle locking AGI and AGF buffers for all scrubber functions. If it detects a conflict between scrub and the running transactions, it will try to wait for intents to complete. If the caller of the helper has not enabled the static key, the helper will return -EDEADLOCK, which should result in the scrub being restarted with the `TRY_HARDER` flag set. The scrub setup function should detect that flag, enable the static key, and try the scrub again. Scrub teardown disables all static keys obtained by `xchk_fsgates_enable`.

For more information, please see the kernel documentation of Documentation/staging/static-keys.rst.

### Pageable Kernel Memory

Some online checking functions work by scanning the filesystem to build a shadow copy of an ondisk metadata structure in memory and comparing the two copies. For online repair to rebuild a metadata structure, it must compute the record set that will be stored in the new structure before it can persist that new structure to disk. Ideally, repairs complete with a single atomic commit that introduces a new data structure. To meet these goals, the kernel needs to collect a large amount of information in a place that doesn't require the correct operation of the filesystem.

Kernel memory isn't suitable because:

- Allocating a contiguous region of memory to create a C array is very difficult, especially on 32-bit systems.
- Linked lists of records introduce double pointer overhead which is very high and eliminate the possibility of indexed lookups.
- Kernel memory is pinned, which can drive the system into OOM conditions.
- The system might not have sufficient memory to stage all the information.

At any given time, online fsck does not need to keep the entire record set in memory, which means that individual records can be paged out if necessary. Continued development of online fsck demonstrated that the ability to perform indexed data storage would also be very useful. Fortunately, the Linux kernel already has a facility for byte-addressable and pageable storage: tmpfs. In-kernel graphics drivers (most notably i915) take advantage of tmpfs files to store intermediate data that doesn't need to be in memory at all times, so that usage precedent is already established. Hence, the `xfile` was born!

| **Historical Sidebar**: |
| --- |
| The first edition of online repair inserted records into a new btree as it found them, which failed because filesystem could shut down with a built data structure, which would be live after recovery finished.<br>The second edition solved the half-rebuilt structure problem by storing everything in memory, but frequently ran the system out of memory.<br>The third edition solved the OOM problem by using linked lists, but the memory overhead of the list pointers was extreme. |

### xfile Access Models

A survey of the intended uses of xfiles suggested these use cases:

1. Arrays of fixed-sized records (space management btrees, directory and extended attribute entries)
2. Sparse arrays of fixed-sized records (quotas and link counts)
3. Large binary objects (BLOBs) of variable sizes (directory and extended attribute names and values)
4. Staging btrees in memory (reverse mapping btrees)
5. Arbitrary contents (realtime space management)

To support the first four use cases, high level data structures wrap the xfile to share functionality between online fsck functions. The rest of this section discusses the interfaces that the xfile presents to four of those five higher level data structures. The fifth use case is discussed in the *realtime summary* case study.

The most general storage interface supported by the xfile enables the reading and writing of arbitrary quantities of data at arbitrary offsets in the xfile. This capability is provided by `xfile_pread` and `xfile_pwrite` functions, which behave similarly to their userspace counterparts. XFS is very record-based, which suggests that the ability to load and store complete records is important. To support these cases, a pair of `xfile_obj_load` and `xfile_obj_store` functions are provided to read and persist objects into an xfile. They are internally the same as pread and pwrite, except that they treat any error as an out of memory error. For online repair, squashing error conditions in this manner is an acceptable behavior because the only reaction is to abort the operation back to userspace. All five xfile usecases can be serviced by these four functions.

However, no discussion of file access idioms is complete without answering the question, ”But what about mmap?” It is convenient to access storage directly with pointers, just like userspace code does with regular memory. Online fsck must not drive the system into OOM conditions, which means that xfiles must be responsive to memory reclamation. tmpfs can only push a pagecache folio to the swap cache if the folio is neither pinned nor locked, which means the xfile must not pin too many folios.

Short term direct access to xfile contents is done by locking the pagecache folio and mapping it into kernel address space. Programmatic access (e.g. pread and pwrite) uses this mechanism. Folio locks are not supposed to be held for long periods of time, so long term direct access to xfile contents is done by bumping the folio refcount, mapping it into kernel address space, and dropping the folio lock. These long term users *must* be responsive to memory reclaim by hooking into the shrinker infrastructure to know when to release folios.

The `xfile_get_page` and `xfile_put_page` functions are provided to retrieve the (locked) folio that backs part of an xfile and to release it. The only code to use these folio lease functions are the xfarray *sorting* algorithms and the *in-memory btrees*.

### xfile Access Coordination

For security reasons, xfiles must be owned privately by the kernel. They are marked `S_PRIVATE` to prevent interference from the security system, must never be mapped into process file descriptor tables, and their pages must never be mapped into userspace processes.

To avoid locking recursion issues with the VFS, all accesses to the shmfs file are performed by manipulating the page cache directly. xfile writers call the `->write_begin` and `->write_end` functions of the xfile's address space to grab writable pages, copy the caller's buffer into the page, and release the pages. xfile readers call `shmem_read_mapping_page_gfp` to grab pages directly before copying the contents into the caller's buffer. In other words, xfiles ignore the VFS read and write code paths to avoid having to create a dummy `struct kiocb` and to avoid taking inode and freeze locks. tmpfs cannot be frozen, and xfiles must not be exposed to userspace.

If an xfile is shared between threads to stage repairs, the caller must provide its own locks to coordinate access. For example, if a scrub function stores scan results in an xfile and needs other threads to provide updates to the scanned data, the scrub function must provide a lock for all threads to share.

### Arrays of Fixed-Sized Records

In XFS, each type of indexed space metadata (free space, inodes, reference counts, file fork space, and reverse mappings) consists of a set of fixed-size records indexed with a classic B+ tree. Directories have a set of fixed-size dirent records that point to the names, and extended attributes have a set of fixed-size attribute keys that point to names and values. Quota counters and file link counters index records with numbers. During a repair, scrub needs to stage new records during the gathering step and retrieve them during the btree building step.

Although this requirement can be satisfied by calling the read and write methods of the xfile directly, it is simpler for callers for there to be a higher level abstraction to take care of computing array offsets, to provide iterator functions, and to deal with sparse records and sorting. The `xfarray` abstraction presents a linear array for fixed-size records atop the byte-accessible xfile.

### Array Access Patterns

Array access patterns in online fsck tend to fall into three categories. Iteration of records is assumed to be necessary for all cases and will be covered in the next section.

The first type of caller handles records that are indexed by position. Gaps may exist between records, and a record may be updated multiple times during the collection step. In other words, these callers want a sparse linearly addressed table file. The typical use case are quota records or file link count records. Access to array elements is performed programmatically via `xfarray_load` and `xfarray_store` functions, which wrap the similarly-named xfile functions to provide loading and storing of array elements at arbitrary array indices. Gaps are defined to be null records, and null records are defined to be a sequence of all zero bytes. Null records are detected by calling `xfarray_element_is_null`. They are created either by calling `xfarray_unset` to null out an existing record or by never storing anything to an array index.

The second type of caller handles records that are not indexed by position and do not require multiple updates to a record. The typical use case here is rebuilding space btrees and key/value btrees. These callers can add records to the array without caring about array indices via the

xfarray_append function, which stores a record at the end of the array. For callers that require records to be presentable in a specific order (e.g. rebuilding btree data), the xfarray_sort function can arrange the sorted records; this function will be covered later.

The third type of caller is a bag, which is useful for counting records. The typical use case here is constructing space extent reference counts from reverse mapping information. Records can be put in the bag in any order, they can be removed from the bag at any time, and uniqueness of records is left to callers. The xfarray_store_anywhere function is used to insert a record in any null record slot in the bag; and the xfarray_unset function removes a record from the bag.

The proposed patchset is the big in-memory array.

### Iterating Array Elements

Most users of the xfarray require the ability to iterate the records stored in the array. Callers can probe every possible array index with the following:

```
xfarray_idx_t i;
foreach_xfarray_idx(array, i) {
    xfarray_load(array, i, &rec);

    /* do something with rec */
}
```

All users of this idiom must be prepared to handle null records or must already know that there aren't any.

For xfarray users that want to iterate a sparse array, the xfarray_iter function ignores indices in the xfarray that have never been written to by calling xfile_seek_data (which internally uses SEEK_DATA) to skip areas of the array that are not populated with memory pages. Once it finds a page, it will skip the zeroed areas of the page.

```
xfarray_idx_t i = XFARRAY_CURSOR_INIT;
while ((ret = xfarray_iter(array, &i, &rec)) == 1) {
    /* do something with rec */
}
```

### Sorting Array Elements

During the fourth demonstration of online repair, a community reviewer remarked that for performance reasons, online repair ought to load batches of records into btree record blocks instead of inserting records into a new btree one at a time. The btree insertion code in XFS is responsible for maintaining correct ordering of the records, so naturally the xfarray must also support sorting the record set prior to bulk loading.

### Case Study: Sorting xfarrays

The sorting algorithm used in the xfarray is actually a combination of adaptive quicksort and a heapsort subalgorithm in the spirit of Sedgewick and pdqsort, with customizations for the Linux kernel. To sort records in a reasonably short amount of time, `xfarray` takes advantage of the binary subpartitioning offered by quicksort, but it also uses heapsort to hedge against performance collapse if the chosen quicksort pivots are poor. Both algorithms are (in general) O(n * lg(n)), but there is a wide performance gulf between the two implementations.

The Linux kernel already contains a reasonably fast implementation of heapsort. It only operates on regular C arrays, which limits the scope of its usefulness. There are two key places where the xfarray uses it:

- Sorting any record subset backed by a single xfile page.
- Loading a small number of xfarray records from potentially disparate parts of the xfarray into a memory buffer, and sorting the buffer.

In other words, `xfarray` uses heapsort to constrain the nested recursion of quicksort, thereby mitigating quicksort's worst runtime behavior.

Choosing a quicksort pivot is a tricky business. A good pivot splits the set to sort in half, leading to the divide and conquer behavior that is crucial to O(n * lg(n)) performance. A poor pivot barely splits the subset at all, leading to $O(n^2)$ runtime. The xfarray sort routine tries to avoid picking a bad pivot by sampling nine records into a memory buffer and using the kernel heapsort to identify the median of the nine.

Most modern quicksort implementations employ Tukey's ”ninther” to select a pivot from a classic C array. Typical ninther implementations pick three unique triads of records, sort each of the triads, and then sort the middle value of each triad to determine the ninther value. As stated previously, however, xfile accesses are not entirely cheap. It turned out to be much more performant to read the nine elements into a memory buffer, run the kernel's in-memory heapsort on the buffer, and choose the 4th element of that buffer as the pivot. Tukey's ninthers are described in J. W. Tukey, *The ninther, a technique for low-effort robust (resistant) location in large samples*, in *Contributions to Survey Sampling and Applied Statistics*, edited by H. David, (Academic Press, 1978), pp. 251–257.

The partitioning of quicksort is fairly textbook -- rearrange the record subset around the pivot, then set up the current and next stack frames to sort with the larger and the smaller halves of the pivot, respectively. This keeps the stack space requirements to log2(record count).

As a final performance optimization, the hi and lo scanning phase of quicksort keeps examined xfile pages mapped in the kernel for as long as possible to reduce map/unmap cycles. Surprisingly, this reduces overall sort runtime by nearly half again after accounting for the application of heapsort directly onto xfile pages.

### Blob Storage

Extended attributes and directories add an additional requirement for staging records: arbitrary byte sequences of finite length. Each directory entry record needs to store entry name, and each extended attribute needs to store both the attribute name and value. The names, keys, and values can consume a large amount of memory, so the `xfblob` abstraction was created to simplify management of these blobs atop an xfile.

Blob arrays provide `xfblob_load` and `xfblob_store` functions to retrieve and persist objects. The store function returns a magic cookie for every object that it persists. Later, callers provide this cookie to the `xblob_load` to recall the object. The `xfblob_free` function frees a specific blob, and the `xfblob_truncate` function frees them all because compaction is not needed.

The details of repairing directories and extended attributes will be discussed in a subsequent section about atomic extent swapping. However, it should be noted that these repair functions only use blob storage to cache a small number of entries before adding them to a temporary ondisk file, which is why compaction is not required.

The proposed patchset is at the start of the extended attribute repair series.

### In-Memory B+Trees

The chapter about *secondary metadata* mentioned that checking and repairing of secondary metadata commonly requires coordination between a live metadata scan of the filesystem and writer threads that are updating that metadata. Keeping the scan data up to date requires requires the ability to propagate metadata updates from the filesystem into the data being collected by the scan. This *can* be done by appending concurrent updates into a separate log file and applying them before writing the new metadata to disk, but this leads to unbounded memory consumption if the rest of the system is very busy. Another option is to skip the sidelog and commit live updates from the filesystem directly into the scan data, which trades more overhead for a lower maximum memory requirement. In both cases, the data structure holding the scan results must support indexed access to perform well.

Given that indexed lookups of scan data is required for both strategies, online fsck employs the second strategy of committing live updates directly into scan data. Because xfarrays are not indexed and do not enforce record ordering, they are not suitable for this task. Conveniently, however, XFS has a library to create and maintain ordered reverse mapping records: the existing rmap btree code! If only there was a means to create one in memory.

Recall that the *xfile* abstraction represents memory pages as a regular file, which means that the kernel can create byte or block addressable virtual address spaces at will. The XFS buffer cache specializes in abstracting IO to block-oriented address spaces, which means that adaptation of the buffer cache to interface with xfiles enables reuse of the entire btree library. Btrees built atop an xfile are collectively known as `xfbtrees`. The next few sections describe how they actually work.

The proposed patchset is the in-memory btree series.

### Using xfiles as a Buffer Cache Target

Two modifications are necessary to support xfiles as a buffer cache target. The first is to make it possible for the `struct xfs_buftarg` structure to host the `struct xfs_buf` rhashtable, because normally those are held by a per-AG structure. The second change is to modify the buffer `ioapply` function to "read" cached pages from the xfile and "write" cached pages back to the xfile. Multiple access to individual buffers is controlled by the `xfs_buf` lock, since the xfile does not provide any locking on its own. With this adaptation in place, users of the xfile-backed buffer cache use exactly the same APIs as users of the disk-backed buffer cache. The separation between xfile and buffer cache implies higher memory usage since they do not share pages, but this property could some day enable transactional updates to an in-memory btree. Today, however, it simply eliminates the need for new code.

### Space Management with an xfbtree

Space management for an xfile is very simple -- each btree block is one memory page in size. These blocks use the same header format as an on-disk btree, but the in-memory block verifiers ignore the checksums, assuming that xfile memory is no more corruption-prone than regular DRAM. Reusing existing code here is more important than absolute memory efficiency.

The very first block of an xfile backing an xfbtree contains a header block. The header describes the owner, height, and the block number of the root xfbtree block.

To allocate a btree block, use `xfile_seek_data` to find a gap in the file. If there are no gaps, create one by extending the length of the xfile. Preallocate space for the block with `xfile_prealloc`, and hand back the location. To free an xfbtree block, use `xfile_discard` (which internally uses `FALLOC_FL_PUNCH_HOLE`) to remove the memory page from the xfile.

### Populating an xfbtree

An online fsck function that wants to create an xfbtree should proceed as follows:

1. Call `xfile_create` to create an xfile.
2. Call `xfs_alloc_memory_buftarg` to create a buffer cache target structure pointing to the xfile.
3. Pass the buffer cache target, buffer ops, and other information to `xfbtree_create` to write an initial tree header and root block to the xfile. Each btree type should define a wrapper that passes necessary arguments to the creation function. For example, rmap btrees define `xfs_rmapbt_mem_create` to take care of all the necessary details for callers. A `struct xfbtree` object will be returned.
4. Pass the xfbtree object to the btree cursor creation function for the btree type. Following the example above, `xfs_rmapbt_mem_cursor` takes care of this for callers.
5. Pass the btree cursor to the regular btree functions to make queries against and to update the in-memory btree. For example, a btree cursor for an rmap xfbtree can be passed to the `xfs_rmap_*` functions just like any other btree cursor. See the *next section* for information on dealing with xfbtree updates that are logged to a transaction.
6. When finished, delete the btree cursor, destroy the xfbtree object, free the buffer target, and the destroy the xfile to release all resources.

### Committing Logged xfbtree Buffers

Although it is a clever hack to reuse the rmap btree code to handle the staging structure, the ephemeral nature of the in-memory btree block storage presents some challenges of its own. The XFS transaction manager must not commit buffer log items for buffers backed by an xfile because the log format does not understand updates for devices other than the data device. An ephemeral xfbtree probably will not exist by the time the AIL checkpoints log transactions back into the filesystem, and certainly won't exist during log recovery. For these reasons, any code updating an xfbtree in transaction context must remove the buffer log items from the transaction and write the updates into the backing xfile before committing or cancelling the transaction.

The `xfbtree_trans_commit` and `xfbtree_trans_cancel` functions implement this functionality as follows:

1. Find each buffer log item whose buffer targets the xfile.
2. Record the dirty/ordered status of the log item.
3. Detach the log item from the buffer.
4. Queue the buffer to a special delwri list.
5. Clear the transaction dirty flag if the only dirty log items were the ones that were detached in step 3.
6. Submit the delwri list to commit the changes to the xfile, if the updates are being committed.

After removing xfile logged buffers from the transaction in this manner, the transaction can be committed or cancelled.

### Bulk Loading of Ondisk B+Trees

As mentioned previously, early iterations of online repair built new btree structures by creating a new btree and adding observations individually. Loading a btree one record at a time had a slight advantage of not requiring the incore records to be sorted prior to commit, but was very slow and leaked blocks if the system went down during a repair. Loading records one at a time also meant that repair could not control the loading factor of the blocks in the new btree.

Fortunately, the venerable `xfs_repair` tool had a more efficient means for rebuilding a btree index from a collection of records -- bulk btree loading. This was implemented rather inefficiently code-wise, since `xfs_repair` had separate copy-pasted implementations for each btree type.

To prepare for online fsck, each of the four bulk loaders were studied, notes were taken, and the four were refactored into a single generic btree bulk loading mechanism. Those notes in turn have been refreshed and are presented below.

### Geometry Computation

The zeroth step of bulk loading is to assemble the entire record set that will be stored in the new btree, and sort the records. Next, call `xfs_btree_bload_compute_geometry` to compute the shape of the btree from the record set, the type of btree, and any load factor preferences. This information is required for resource reservation.

First, the geometry computation computes the minimum and maximum records that will fit in a leaf block from the size of a btree block and the size of the block header. Roughly speaking, the maximum number of records is:

```
maxrecs = (block_size - header_size) / record_size
```

The XFS design specifies that btree blocks should be merged when possible, which means the minimum number of records is half of maxrecs:

```
minrecs = maxrecs / 2
```

The next variable to determine is the desired loading factor. This must be at least minrecs and no more than maxrecs. Choosing minrecs is undesirable because it wastes half the block. Choosing maxrecs is also undesirable because adding a single record to each newly rebuilt leaf block will cause a tree split, which causes a noticeable drop in performance immediately afterwards. The default loading factor was chosen to be 75% of maxrecs, which provides a reasonably compact structure without any immediate split penalties:

```
default_load_factor = (maxrecs + minrecs) / 2
```

If space is tight, the loading factor will be set to maxrecs to try to avoid running out of space:

```
leaf_load_factor = enough space ? default_load_factor : maxrecs
```

Load factor is computed for btree node blocks using the combined size of the btree key and pointer as the record size:

```
maxrecs = (block_size - header_size) / (key_size + ptr_size)
minrecs = maxrecs / 2
node_load_factor = enough space ? default_load_factor : maxrecs
```

Once that's done, the number of leaf blocks required to store the record set can be computed as:

```
leaf_blocks = ceil(record_count / leaf_load_factor)
```

The number of node blocks needed to point to the next level down in the tree is computed as:

```
n_blocks = (n == 0 ? leaf_blocks : node_blocks[n])
node_blocks[n + 1] = ceil(n_blocks / node_load_factor)
```

The entire computation is performed recursively until the current level only needs one block. The resulting geometry is as follows:

- For AG-rooted btrees, this level is the root level, so the height of the new tree is `level + 1` and the space needed is the summation of the number of blocks on each level.

- For inode-rooted btrees where the records in the top level do not fit in the inode fork area, the height is `level + 2`, the space needed is the summation of the number of blocks on each level, and the inode fork points to the root block.
- For inode-rooted btrees where the records in the top level can be stored in the inode fork area, then the root block can be stored in the inode, the height is `level + 1`, and the space needed is one less than the summation of the number of blocks on each level. This only becomes relevant when non-bmap btrees gain the ability to root in an inode, which is a future patchset and only included here for completeness.

### Reserving New B+Tree Blocks

Once repair knows the number of blocks needed for the new btree, it allocates those blocks using the free space information. Each reserved extent is tracked separately by the btree builder state data. To improve crash resilience, the reservation code also logs an Extent Freeing Intent (EFI) item in the same transaction as each space allocation and attaches its in-memory `struct xfs_extent_free_item` object to the space reservation. If the system goes down, log recovery will use the unfinished EFIs to free the unused space, the free space, leaving the filesystem unchanged.

Each time the btree builder claims a block for the btree from a reserved extent, it updates the in-memory reservation to reflect the claimed space. Block reservation tries to allocate as much contiguous space as possible to reduce the number of EFIs in play.

While repair is writing these new btree blocks, the EFIs created for the space reservations pin the tail of the ondisk log. It's possible that other parts of the system will remain busy and push the head of the log towards the pinned tail. To avoid livelocking the filesystem, the EFIs must not pin the tail of the log for too long. To alleviate this problem, the dynamic relogging capability of the deferred ops mechanism is reused here to commit a transaction at the log head containing an EFD for the old EFI and new EFI at the head. This enables the log to release the old EFI to keep the log moving forwards.

EFIs have a role to play during the commit and reaping phases; please see the next section and the section about *reaping* for more details.

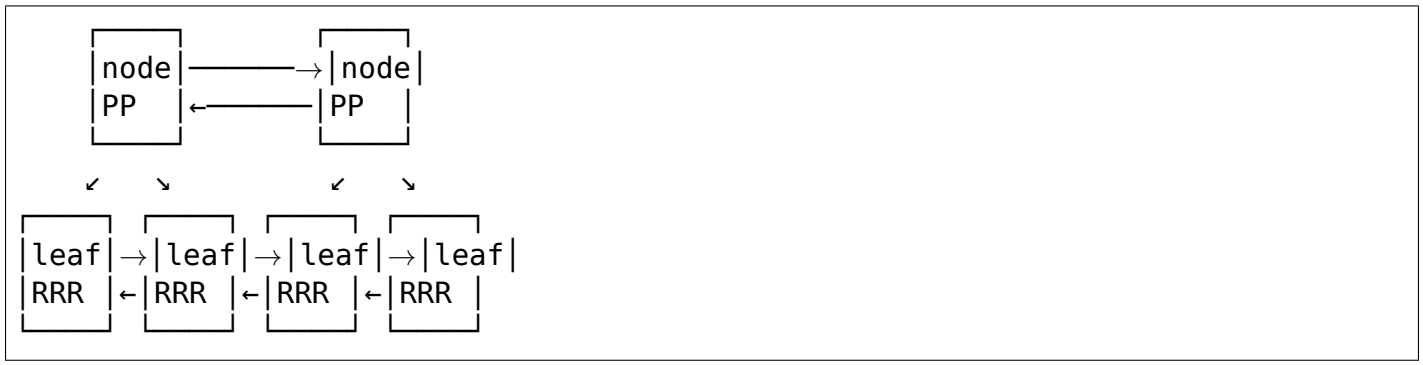
Proposed patchsets are the bitmap rework and the preparation for bulk loading btrees.

### Writing the New Tree

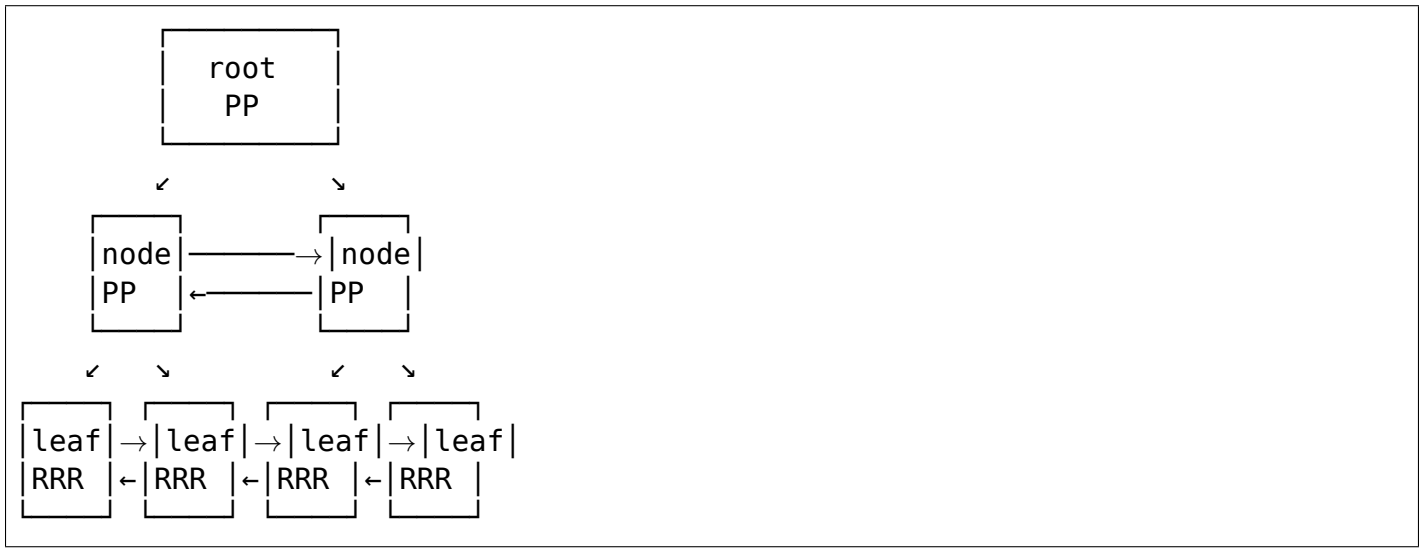
This part is pretty simple -- the btree builder (`xfs_btree_bulkload`) claims a block from the reserved list, writes the new btree block header, fills the rest of the block with records, and adds the new leaf block to a list of written blocks:


Sibling pointers are set every time a new block is added to the level:


When it finishes writing the record leaf blocks, it moves on to the node blocks To fill a node block, it walks each block in the next level down in the tree to compute the relevant keys and write them into the parent node:




When it reaches the root level, it is ready to commit the new btree!:




The first step to commit the new btree is to persist the btree blocks to disk synchronously. This is a little complicated because a new btree block could have been freed in the recent past, so the builder must use `xfs_buf_delwri_queue_here` to remove the (stale) buffer from the AIL list before it can write the new blocks to disk. Blocks are queued for IO using a delwri list and written in one large batch with `xfs_buf_delwri_submit`.

Once the new blocks have been persisted to disk, control returns to the individual repair function that called the bulk loader. The repair function must log the location of the new root in a transaction, clean up the space reservations that were made for the new btree, and reap the old metadata blocks:

1. Commit the location of the new btree root.
2. For each incore reservation:
    a. Log Extent Freeing Done (EFD) items for all the space that was consumed by the btree builder. The new EFDs must point to the EFIs attached to the reservation to prevent log recovery from freeing the new blocks.

   b. For unclaimed portions of incore reservations, create a regular deferred extent free work item to be free the unused space later in the transaction chain.

   c. The EFDs and EFIs logged in steps 2a and 2b must not overrun the reservation of the committing transaction. If the btree loading code suspects this might be about to happen, it must call `xrep_defer_finish` to clear out the deferred work and obtain a fresh transaction.

3. Clear out the deferred work a second time to finish the commit and clean the repair transaction.

The transaction rolling in steps 2c and 3 represent a weakness in the repair algorithm, because a log flush and a crash before the end of the reap step can result in space leaking. Online repair functions minimize the chances of this occurring by using very large transactions, which each can accommodate many thousands of block freeing instructions. Repair moves on to reaping the old blocks, which will be presented in a subsequent *section* after a few case studies of bulk loading.

### Case Study: Rebuilding the Inode Index

The high level process to rebuild the inode index btree is:

1. Walk the reverse mapping records to generate `struct xfs_inobt_rec` records from the inode chunk information and a bitmap of the old inode btree blocks.
2. Append the records to an xfarray in inode order.
3. Use the `xfs_btree_bload_compute_geometry` function to compute the number of blocks needed for the inode btree. If the free space inode btree is enabled, call it again to estimate the geometry of the finobt.
4. Allocate the number of blocks computed in the previous step.
5. Use `xfs_btree_bload` to write the xfarray records to btree blocks and generate the internal node blocks. If the free space inode btree is enabled, call it again to load the finobt.
6. Commit the location of the new btree root block(s) to the AGI.
7. Reap the old btree blocks using the bitmap created in step 1.

Details are as follows.

The inode btree maps inumbers to the ondisk location of the associated inode records, which means that the inode btrees can be rebuilt from the reverse mapping information. Reverse mapping records with an owner of `XFS_RMAP_OWN_INOBT` marks the location of the old inode btree blocks. Each reverse mapping record with an owner of `XFS_RMAP_OWN_INODES` marks the location of at least one inode cluster buffer. A cluster is the smallest number of ondisk inodes that can be allocated or freed in a single transaction; it is never smaller than 1 fs block or 4 inodes.

For the space represented by each inode cluster, ensure that there are no records in the free space btrees nor any records in the reference count btree. If there are, the space metadata inconsistencies are reason enough to abort the operation. Otherwise, read each cluster buffer to check that its contents appear to be ondisk inodes and to decide if the file is allocated (`xfs_dinode.i_mode != 0`) or free (`xfs_dinode.i_mode == 0`). Accumulate the results of successive inode cluster buffer reads until there is enough information to fill a single inode chunk

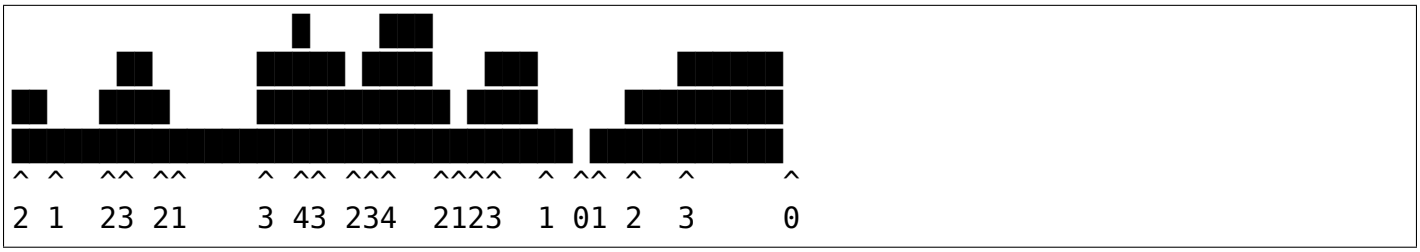
record, which is 64 consecutive numbers in the inumber keyspace. If the chunk is sparse, the chunk record may include holes.

Once the repair function accumulates one chunk's worth of data, it calls `xfarray_append` to add the inode btree record to the xfarray. This xfarray is walked twice during the btree creation step -- once to populate the inode btree with all inode chunk records, and a second time to populate the free inode btree with records for chunks that have free non-sparse inodes. The number of records for the inode btree is the number of xfarray records, but the record count for the free inode btree has to be computed as inode chunk records are stored in the xfarray.

The proposed patchset is the AG btree repair series.

### Case Study: Rebuilding the Space Reference Counts

Reverse mapping records are used to rebuild the reference count information. Reference counts are required for correct operation of copy on write for shared file data. Imagine the reverse mapping entries as rectangles representing extents of physical blocks, and that the rectangles can be laid down to allow them to overlap each other. From the diagram below, it is apparent that a reference count record must start or end wherever the height of the stack changes. In other words, the record emission stimulus is level-triggered:




The ondisk reference count btree does not store the refcount == 0 cases because the free space btree already records which blocks are free. Extents being used to stage copy-on-write operations should be the only records with refcount == 1. Single-owner file blocks aren't recorded in either the free space or the reference count btrees.

The high level process to rebuild the reference count btree is:

1. Walk the reverse mapping records to generate `struct xfs_refcount_irec` records for any space having more than one reverse mapping and add them to the xfarray. Any records owned by `XFS_RMAP_OWN_COW` are also added to the xfarray because these are extents allocated to stage a copy on write operation and are tracked in the refcount btree.

   Use any records owned by `XFS_RMAP_OWN_REFC` to create a bitmap of old refcount btree blocks.

2. Sort the records in physical extent order, putting the CoW staging extents at the end of the xfarray. This matches the sorting order of records in the refcount btree.

3. Use the `xfs_btree_bload_compute_geometry` function to compute the number of blocks needed for the new tree.

4. Allocate the number of blocks computed in the previous step.

5. Use `xfs_btree_bload` to write the xfarray records to btree blocks and generate the internal node blocks.

6. Commit the location of new btree root block to the AGF.

7. Reap the old btree blocks using the bitmap created in step 1.

Details are as follows; the same algorithm is used by `xfs_repair` to generate refcount information from reverse mapping records.

- Until the reverse mapping btree runs out of records:
  - Retrieve the next record from the btree and put it in a bag.
  - Collect all records with the same starting block from the btree and put them in the bag.
  - While the bag isn't empty:
    * Among the mappings in the bag, compute the lowest block number where the reference count changes. This position will be either the starting block number of the next unprocessed reverse mapping or the next block after the shortest mapping in the bag.
    * Remove all mappings from the bag that end at this position.
    * Collect all reverse mappings that start at this position from the btree and put them in the bag.
    * If the size of the bag changed and is greater than one, create a new refcount record associating the block number range that we just walked to the size of the bag.

The bag-like structure in this case is a type 2 xfarray as discussed in the *xfarray access patterns* section. Reverse mappings are added to the bag using `xfarray_store_anywhere` and removed via `xfarray_unset`. Bag members are examined through `xfarray_iter` loops.

The proposed patchset is the AG btree repair series.

#### Case Study: Rebuilding File Fork Mapping Indices

The high level process to rebuild a data/attr fork mapping btree is:

1. Walk the reverse mapping records to generate `struct xfs_bmbt_rec` records from the reverse mapping records for that inode and fork. Append these records to an xfarray. Compute the bitmap of the old bmap btree blocks from the `BMBT_BLOCK` records.
2. Use the `xfs_btree_bload_compute_geometry` function to compute the number of blocks needed for the new tree.
3. Sort the records in file offset order.
4. If the extent records would fit in the inode fork immediate area, commit the records to that immediate area and skip to step 8.
5. Allocate the number of blocks computed in the previous step.
6. Use `xfs_btree_bload` to write the xfarray records to btree blocks and generate the internal node blocks.
7. Commit the new btree root block to the inode fork immediate area.
8. Reap the old btree blocks using the bitmap created in step 1.

There are some complications here: First, it's possible to move the fork offset to adjust the sizes of the immediate areas if the data and attr forks are not both in BMBT format. Second, if there

are sufficiently few fork mappings, it may be possible to use EXTENTS format instead of BMBT, which may require a conversion. Third, the incore extent map must be reloaded carefully to avoid disturbing any delayed allocation extents.

The proposed patchset is the file mapping repair series.

### Reaping Old Metadata Blocks

Whenever online fsck builds a new data structure to replace one that is suspect, there is a question of how to find and dispose of the blocks that belonged to the old structure. The laziest method of course is not to deal with them at all, but this slowly leads to service degradations as space leaks out of the filesystem. Hopefully, someone will schedule a rebuild of the free space information to plug all those leaks. Offline repair rebuilds all space metadata after recording the usage of the files and directories that it decides not to clear, hence it can build new structures in the discovered free space and avoid the question of reaping.

As part of a repair, online fsck relies heavily on the reverse mapping records to find space that is owned by the corresponding rmap owner yet truly free. Cross referencing rmap records with other rmap records is necessary because there may be other data structures that also think they own some of those blocks (e.g. crosslinked trees). Permitting the block allocator to hand them out again will not push the system towards consistency.

For space metadata, the process of finding extents to dispose of generally follows this format:

1. Create a bitmap of space used by data structures that must be preserved. The space reservations used to create the new metadata can be used here if the same rmap owner code is used to denote all of the objects being rebuilt.
2. Survey the reverse mapping data to create a bitmap of space owned by the same `XFS_RMAP_OWN_*` number for the metadata that is being preserved.
3. Use the bitmap disunion operator to subtract (1) from (2). The remaining set bits represent candidate extents that could be freed. The process moves on to step 4 below.

Repairs for file-based metadata such as extended attributes, directories, symbolic links, quota files and realtime bitmaps are performed by building a new structure attached to a temporary file and swapping the forks. Afterward, the mappings in the old file fork are the candidate blocks for disposal.

The process for disposing of old extents is as follows:

4. For each candidate extent, count the number of reverse mapping records for the first block in that extent that do not have the same rmap owner for the data structure being repaired.
   - If zero, the block has a single owner and can be freed.
   - If not, the block is part of a crosslinked structure and must not be freed.
5. Starting with the next block in the extent, figure out how many more blocks have the same zero/nonzero other owner status as that first block.
6. If the region is crosslinked, delete the reverse mapping entry for the structure being repaired and move on to the next region.
7. If the region is to be freed, mark any corresponding buffers in the buffer cache as stale to prevent log writeback.
8. Free the region and move on.

However, there is one complication to this procedure. Transactions are of finite size, so the reaping process must be careful to roll the transactions to avoid overruns. Overruns come from two sources:

a. EFIs logged on behalf of space that is no longer occupied

b. Log items for buffer invalidations

This is also a window in which a crash during the reaping process can leak blocks. As stated earlier, online repair functions use very large transactions to minimize the chances of this occurring.

The proposed patchset is the preparation for bulk loading btrees series.

### Case Study: Reaping After a Regular Btree Repair

Old reference count and inode btrees are the easiest to reap because they have rmap records with special owner codes: `XFS_RMAP_OWN_REFC` for the refcount btree, and `XFS_RMAP_OWN_INOBT` for the inode and free inode btrees. Creating a list of extents to reap the old btree blocks is quite simple, conceptually:

1. Lock the relevant AGI/AGF header buffers to prevent allocation and frees.
2. For each reverse mapping record with an rmap owner corresponding to the metadata structure being rebuilt, set the corresponding range in a bitmap.
3. Walk the current data structures that have the same rmap owner. For each block visited, clear that range in the above bitmap.
4. Each set bit in the bitmap represents a block that could be a block from the old data structures and hence is a candidate for reaping. In other words, `(rmap_records_owned_by & ~blocks_reachable_by_walk)` are the blocks that might be freeable.

If it is possible to maintain the AGF lock throughout the repair (which is the common case), then step 2 can be performed at the same time as the reverse mapping record walk that creates the records for the new btree.

### Case Study: Rebuilding the Free Space Indices

The high level process to rebuild the free space indices is:

1. Walk the reverse mapping records to generate `struct xfs_alloc_rec_incore` records from the gaps in the reverse mapping btree.
2. Append the records to an xfarray.
3. Use the `xfs_btree_bload_compute_geometry` function to compute the number of blocks needed for each new tree.
4. Allocate the number of blocks computed in the previous step from the free space information collected.
5. Use `xfs_btree_bload` to write the xfarray records to btree blocks and generate the internal node blocks for the free space by length index. Call it again for the free space by block number index.
6. Commit the locations of the new btree root blocks to the AGF.

7. Reap the old btree blocks by looking for space that is not recorded by the reverse mapping btree, the new free space btrees, or the AGFL.

Repairing the free space btrees has three key complications over a regular btree repair:

First, free space is not explicitly tracked in the reverse mapping records. Hence, the new free space records must be inferred from gaps in the physical space component of the keyspace of the reverse mapping btree.

Second, free space repairs cannot use the common btree reservation code because new blocks are reserved out of the free space btrees. This is impossible when repairing the free space btrees themselves. However, repair holds the AGF buffer lock for the duration of the free space index reconstruction, so it can use the collected free space information to supply the blocks for the new free space btrees. It is not necessary to back each reserved extent with an EFI because the new free space btrees are constructed in what the ondisk filesystem thinks is unowned space. However, if reserving blocks for the new btrees from the collected free space information changes the number of free space records, repair must re-estimate the new free space btree geometry with the new record count until the reservation is sufficient. As part of committing the new btrees, repair must ensure that reverse mappings are created for the reserved blocks and that unused reserved blocks are inserted into the free space btrees. Deferrred rmap and freeing operations are used to ensure that this transition is atomic, similar to the other btree repair functions.

Third, finding the blocks to reap after the repair is not overly straightforward. Blocks for the free space btrees and the reverse mapping btrees are supplied by the AGFL. Blocks put onto the AGFL have reverse mapping records with the owner `XFS_RMAP_OWN_AG`. This ownership is retained when blocks move from the AGFL into the free space btrees or the reverse mapping btrees. When repair walks reverse mapping records to synthesize free space records, it creates a bitmap (`ag_owner_bitmap`) of all the space claimed by `XFS_RMAP_OWN_AG` records. The repair context maintains a second bitmap corresponding to the rmap btree blocks and the AGFL blocks (`rmap_agfl_bitmap`). When the walk is complete, the bitmap disunion operation (`ag_owner_bitmap & ~rmap_agfl_bitmap`) computes the extents that are used by the old free space btrees. These blocks can then be reaped using the methods outlined above.

The proposed patchset is the AG btree repair series.

### Case Study: Reaping After Repairing Reverse Mapping Btrees

Old reverse mapping btrees are less difficult to reap after a repair. As mentioned in the previous section, blocks on the AGFL, the two free space btree blocks, and the reverse mapping btree blocks all have reverse mapping records with `XFS_RMAP_OWN_AG` as the owner. The full process of gathering reverse mapping records and building a new btree are described in the case study of *live rebuilds of rmap data*, but a crucial point from that discussion is that the new rmap btree will not contain any records for the old rmap btree, nor will the old btree blocks be tracked in the free space btrees. The list of candidate reaping blocks is computed by setting the bits corresponding to the gaps in the new rmap btree records, and then clearing the bits corresponding to extents in the free space btrees and the current AGFL blocks. The result (`new_rmapbt_gaps & ~(agfl | bnobt_records)`) are reaped using the methods outlined above.

The rest of the process of rebuildng the reverse mapping btree is discussed in a separate *case study*.

The proposed patchset is the AG btree repair series.

### Case Study: Rebuilding the AGFL

The allocation group free block list (AGFL) is repaired as follows:

1. Create a bitmap for all the space that the reverse mapping data claims is owned by `XFS_RMAP_OWN_AG`.
2. Subtract the space used by the two free space btrees and the rmap btree.
3. Subtract any space that the reverse mapping data claims is owned by any other owner, to avoid re-adding crosslinked blocks to the AGFL.
4. Once the AGFL is full, reap any blocks leftover.
5. The next operation to fix the freelist will right-size the list.

See fs/xfs/scrub/agheader_repair.c for more details.

### Inode Record Repairs

Inode records must be handled carefully, because they have both ondisk records (”dinodes”) and an in-memory (”cached”) representation. There is a very high potential for cache coherency issues if online fsck is not careful to access the ondisk metadata *only* when the ondisk metadata is so badly damaged that the filesystem cannot load the in-memory representation. When online fsck wants to open a damaged file for scrubbing, it must use specialized resource acquisition functions that return either the in-memory representation *or* a lock on whichever object is necessary to prevent any update to the ondisk location.

The only repairs that should be made to the ondisk inode buffers are whatever is necessary to get the in-core structure loaded. This means fixing whatever is caught by the inode cluster buffer and inode fork verifiers, and retrying the `iget` operation. If the second `iget` fails, the repair has failed.

Once the in-memory representation is loaded, repair can lock the inode and can subject it to comprehensive checks, repairs, and optimizations. Most inode attributes are easy to check and constrain, or are user-controlled arbitrary bit patterns; these are both easy to fix. Dealing with the data and attr fork extent counts and the file block counts is more complicated, because computing the correct value requires traversing the forks, or if that fails, leaving the fields invalid and waiting for the fork fsck functions to run.

The proposed patchset is the inode repair series.

### Quota Record Repairs

Similar to inodes, quota records (”dquots”) also have both ondisk records and an in-memory representation, and hence are subject to the same cache coherency issues. Somewhat confusingly, both are known as dquots in the XFS codebase.

The only repairs that should be made to the ondisk quota record buffers are whatever is necessary to get the in-core structure loaded. Once the in-memory representation is loaded, the only attributes needing checking are obviously bad limits and timer values.

Quota usage counters are checked, repaired, and discussed separately in the section about *live quotacheck*.

The proposed patchset is the quota repair series.

### Freezing to Fix Summary Counters

Filesystem summary counters track availability of filesystem resources such as free blocks, free inodes, and allocated inodes. This information could be compiled by walking the free space and inode indexes, but this is a slow process, so XFS maintains a copy in the ondisk superblock that should reflect the ondisk metadata, at least when the filesystem has been unmounted cleanly. For performance reasons, XFS also maintains incore copies of those counters, which are key to enabling resource reservations for active transactions. Writer threads reserve the worst-case quantities of resources from the incore counter and give back whatever they don't use at commit time. It is therefore only necessary to serialize on the superblock when the superblock is being committed to disk.

The lazy superblock counter feature introduced in XFS v5 took this even further by training log recovery to recompute the summary counters from the AG headers, which eliminated the need for most transactions even to touch the superblock. The only time XFS commits the summary counters is at filesystem unmount. To reduce contention even further, the incore counter is implemented as a percpu counter, which means that each CPU is allocated a batch of blocks from a global incore counter and can satisfy small allocations from the local batch.

The high-performance nature of the summary counters makes it difficult for online fsck to check them, since there is no way to quiesce a percpu counter while the system is running. Although online fsck can read the filesystem metadata to compute the correct values of the summary counters, there's no way to hold the value of a percpu counter stable, so it's quite possible that the counter will be out of date by the time the walk is complete. Earlier versions of online scrub would return to userspace with an incomplete scan flag, but this is not a satisfying outcome for a system administrator. For repairs, the in-memory counters must be stabilized while walking the filesystem metadata to get an accurate reading and install it in the percpu counter.

To satisfy this requirement, online fsck must prevent other programs in the system from initiating new writes to the filesystem, it must disable background garbage collection threads, and it must wait for existing writer programs to exit the kernel. Once that has been established, scrub can walk the AG free space indexes, the inode btrees, and the realtime bitmap to compute the correct value of all four summary counters. This is very similar to a filesystem freeze, though not all of the pieces are necessary:

- The final freeze state is set one higher than `SB_FREEZE_COMPLETE` to prevent other threads from thawing the filesystem, or other scrub threads from initiating another fscounters freeze.
- It does not quiesce the log.

With this code in place, it is now possible to pause the filesystem for just long enough to check and correct the summary counters.

**Historical Sidebar**:

The initial implementation used the actual VFS filesystem freeze mechanism to quiesce filesystem activity. With the filesystem frozen, it is possible to resolve the counter values with exact precision, but there are many problems with calling the VFS methods directly:

- Other programs can unfreeze the filesystem without our knowledge. This leads to incorrect scan results and incorrect repairs.
- Adding an extra lock to prevent others from thawing the filesystem required the addition of a `->freeze_super` function to wrap `freeze_fs()`. This in turn caused other subtle problems because it turns out that the VFS `freeze_super` and `thaw_super` functions can drop the last reference to the VFS superblock, and any subsequent access becomes a UAF bug! This can happen if the filesystem is unmounted while the underlying block device has frozen the filesystem. This problem could be solved by grabbing extra references to the superblock, but it felt suboptimal given the other inadequacies of this approach.
- The log need not be quiesced to check the summary counters, but a VFS freeze initiates one anyway. This adds unnecessary runtime to live fscounter fsck operations.
- Quiescing the log means that XFS flushes the (possibly incorrect) counters to disk as part of cleaning the log.
- A bug in the VFS meant that freeze could complete even when sync_filesystem fails to flush the filesystem and returns an error. This bug was fixed in Linux 5.17.

The proposed patchset is the summary counter cleanup series.

### Full Filesystem Scans

Certain types of metadata can only be checked by walking every file in the entire filesystem to record observations and comparing the observations against what's recorded on disk. Like every other type of online repair, repairs are made by writing those observations to disk in a replacement structure and committing it atomically. However, it is not practical to shut down the entire filesystem to examine hundreds of billions of files because the downtime would be excessive. Therefore, online fsck must build the infrastructure to manage a live scan of all the files in the filesystem. There are two questions that need to be solved to perform a live walk:

- How does scrub manage the scan while it is collecting data?
- How does the scan keep abreast of changes being made to the system by other threads?

### Coordinated Inode Scans

In the original Unix filesystems of the 1970s, each directory entry contained an index number (*inumber*) which was used as an index into on ondisk array (*itable*) of fixed-size records (*inodes*) describing a file's attributes and its data block mapping. This system is described by J. Lions, ”inode (5659)” in *Lions' Commentary on UNIX, 6th Edition*, (Dept. of Computer Science, the University of New South Wales, November 1977), pp. 18-2; and later by D. Ritchie and K. Thompson, ”Implementation of the File System”, from *The UNIX Time-Sharing System*, (The Bell System Technical Journal, July 1978), pp. 1913-4.

XFS retains most of this design, except now inumbers are search keys over all the space in the data section filesystem. They form a continuous keyspace that can be expressed as a 64-bit

integer, though the inodes themselves are sparsely distributed within the keyspace. Scans proceed in a linear fashion across the inumber keyspace, starting from `0x0` and ending at `0xFFFFFFFFFFFFFFFF`. Naturally, a scan through a keyspace requires a scan cursor object to track the scan progress. Because this keyspace is sparse, this cursor contains two parts. The first part of this scan cursor object tracks the inode that will be examined next; call this the examination cursor. Somewhat less obviously, the scan cursor object must also track which parts of the keyspace have already been visited, which is critical for deciding if a concurrent filesystem update needs to be incorporated into the scan data. Call this the visited inode cursor.

Advancing the scan cursor is a multi-step process encapsulated in `xchk_iscan_iter`:

1. Lock the AGI buffer of the AG containing the inode pointed to by the visited inode cursor. This guarantee that inodes in this AG cannot be allocated or freed while advancing the cursor.
2. Use the per-AG inode btree to look up the next inumber after the one that was just visited, since it may not be keyspace adjacent.
3. If there are no more inodes left in this AG:
    a. Move the examination cursor to the point of the inumber keyspace that corresponds to the start of the next AG.
    b. Adjust the visited inode cursor to indicate that it has ”visited” the last possible inode in the current AG's inode keyspace. XFS inumbers are segmented, so the cursor needs to be marked as having visited the entire keyspace up to just before the start of the next AG's inode keyspace.
    c. Unlock the AGI and return to step 1 if there are unexamined AGs in the filesystem.
    d. If there are no more AGs to examine, set both cursors to the end of the inumber keyspace. The scan is now complete.
4. Otherwise, there is at least one more inode to scan in this AG:
    a. Move the examination cursor ahead to the next inode marked as allocated by the inode btree.
    b. Adjust the visited inode cursor to point to the inode just prior to where the examination cursor is now. Because the scanner holds the AGI buffer lock, no inodes could have been created in the part of the inode keyspace that the visited inode cursor just advanced.
5. Get the incore inode for the inumber of the examination cursor. By maintaining the AGI buffer lock until this point, the scanner knows that it was safe to advance the examination cursor across the entire keyspace, and that it has stabilized this next inode so that it cannot disappear from the filesystem until the scan releases the incore inode.
6. Drop the AGI lock and return the incore inode to the caller.

Online fsck functions scan all files in the filesystem as follows:

1. Start a scan by calling `xchk_iscan_start`.
2. Advance the scan cursor (`xchk_iscan_iter`) to get the next inode. If one is provided:
    a. Lock the inode to prevent updates during the scan.
    b. Scan the inode.

c. While still holding the inode lock, adjust the visited inode cursor (`xchk_iscan_mark_visited`) to point to this inode.

d. Unlock and release the inode.

8. Call `xchk_iscan_teardown` to complete the scan.

There are subtleties with the inode cache that complicate grabbing the incore inode for the caller. Obviously, it is an absolute requirement that the inode metadata be consistent enough to load it into the inode cache. Second, if the incore inode is stuck in some intermediate state, the scan coordinator must release the AGI and push the main filesystem to get the inode back into a loadable state.

The proposed patches are the inode scanner series. The first user of the new functionality is the online quotacheck series.

### Inode Management

In regular filesystem code, references to allocated XFS incore inodes are always obtained (`xfs_iget`) outside of transaction context because the creation of the incore context for an existing file does not require metadata updates. However, it is important to note that references to incore inodes obtained as part of file creation must be performed in transaction context because the filesystem must ensure the atomicity of the ondisk inode btree index updates and the initialization of the actual ondisk inode.

References to incore inodes are always released (`xfs_irele`) outside of transaction context because there are a handful of activities that might require ondisk updates:

- The VFS may decide to kick off writeback as part of a `DONTCACHE` inode release.
- Speculative preallocations need to be unreserved.
- An unlinked file may have lost its last reference, in which case the entire file must be inactivated, which involves releasing all of its resources in the ondisk metadata and freeing the inode.

These activities are collectively called inode inactivation. Inactivation has two parts -- the VFS part, which initiates writeback on all dirty file pages, and the XFS part, which cleans up XFS-specific information and frees the inode if it was unlinked. If the inode is unlinked (or unconnected after a file handle operation), the kernel drops the inode into the inactivation machinery immediately.

During normal operation, resource acquisition for an update follows this order to avoid deadlocks:

1. Inode reference (`iget`).
2. Filesystem freeze protection, if repairing (`mnt_want_write_file`).
3. Inode `IOLOCK` (VFS `i_rwsem`) lock to control file IO.
4. Inode `MMAPLOCK` (page cache `invalidate_lock`) lock for operations that can update page cache mappings.
5. Log feature enablement.
6. Transaction log space grant.
7. Space on the data and realtime devices for the transaction.

8. Incore dquot references, if a file is being repaired. Note that they are not locked, merely acquired.
9. Inode `ILOCK` for file metadata updates.
10. AG header buffer locks / Realtime metadata inode ILOCK.
11. Realtime metadata buffer locks, if applicable.
12. Extent mapping btree blocks, if applicable.

Resources are often released in the reverse order, though this is not required. However, online fsck differs from regular XFS operations because it may examine an object that normally is acquired in a later stage of the locking order, and then decide to cross-reference the object with an object that is acquired earlier in the order. The next few sections detail the specific ways in which online fsck takes care to avoid deadlocks.

### iget and irele During a Scrub

An inode scan performed on behalf of a scrub operation runs in transaction context, and possibly with resources already locked and bound to it. This isn't much of a problem for `iget` since it can operate in the context of an existing transaction, as long as all of the bound resources are acquired before the inode reference in the regular filesystem.

When the VFS `iput` function is given a linked inode with no other references, it normally puts the inode on an LRU list in the hope that it can save time if another process re-opens the file before the system runs out of memory and frees it. Filesystem callers can short-circuit the LRU process by setting a `DONTCACHE` flag on the inode to cause the kernel to try to drop the inode into the inactivation machinery immediately.

In the past, inactivation was always done from the process that dropped the inode, which was a problem for scrub because scrub may already hold a transaction, and XFS does not support nesting transactions. On the other hand, if there is no scrub transaction, it is desirable to drop otherwise unused inodes immediately to avoid polluting caches. To capture these nuances, the online fsck code has a separate `xchk_irele` function to set or clear the `DONTCACHE` flag to get the required release behavior.

Proposed patchsets include fixing scrub iget usage and dir iget usage.

### Locking Inodes

In regular filesystem code, the VFS and XFS will acquire multiple IOLOCK locks in a well-known order: parent → child when updating the directory tree, and in numerical order of the addresses of their `struct inode` object otherwise. For regular files, the MMAPLOCK can be acquired after the IOLOCK to stop page faults. If two MMAPLOCKs must be acquired, they are acquired in numerical order of the addresses of their `struct address_space` objects. Due to the structure of existing filesystem code, IOLOCKs and MMAPLOCKs must be acquired before transactions are allocated. If two ILOCKs must be acquired, they are acquired in inumber order.

Inode lock acquisition must be done carefully during a coordinated inode scan. Online fsck cannot abide these conventions, because for a directory tree scanner, the scrub process holds the IOLOCK of the file being scanned and it needs to take the IOLOCK of the file at the other end of the directory link. If the directory tree is corrupt because it contains a cycle, `xfs_scrub` cannot use the regular inode locking functions and avoid becoming trapped in an ABBA deadlock.

Solving both of these problems is straightforward -- any time online fsck needs to take a second lock of the same class, it uses trylock to avoid an ABBA deadlock. If the trylock fails, scrub drops all inode locks and use trylock loops to (re)acquire all necessary resources. Trylock loops enable scrub to check for pending fatal signals, which is how scrub avoids deadlocking the filesystem or becoming an unresponsive process. However, trylock loops means that online fsck must be prepared to measure the resource being scrubbed before and after the lock cycle to detect changes and react accordingly.

### Case Study: Finding a Directory Parent

Consider the directory parent pointer repair code as an example. Online fsck must verify that the dotdot dirent of a directory points up to a parent directory, and that the parent directory contains exactly one dirent pointing down to the child directory. Fully validating this relationship (and repairing it if possible) requires a walk of every directory on the filesystem while holding the child locked, and while updates to the directory tree are being made. The coordinated inode scan provides a way to walk the filesystem without the possibility of missing an inode. The child directory is kept locked to prevent updates to the dotdot dirent, but if the scanner fails to lock a parent, it can drop and relock both the child and the prospective parent. If the dotdot entry changes while the directory is unlocked, then a move or rename operation must have changed the child's parentage, and the scan can exit early.

The proposed patchset is the directory repair series.

### Filesystem Hooks

The second piece of support that online fsck functions need during a full filesystem scan is the ability to stay informed about updates being made by other threads in the filesystem, since comparisons against the past are useless in a dynamic environment. Two pieces of Linux kernel infrastructure enable online fsck to monitor regular filesystem operations: filesystem hooks and *static keys*.

Filesystem hooks convey information about an ongoing filesystem operation to a downstream consumer. In this case, the downstream consumer is always an online fsck function. Because multiple fsck functions can run in parallel, online fsck uses the Linux notifier call chain facility to dispatch updates to any number of interested fsck processes. Call chains are a dynamic list, which means that they can be configured at run time. Because these hooks are private to the XFS module, the information passed along contains exactly what the checking function needs to update its observations.

The current implementation of XFS hooks uses SRCU notifier chains to reduce the impact to highly threaded workloads. Regular blocking notifier chains use a rwsem and seem to have a much lower overhead for single-threaded applications. However, it may turn out that the combination of blocking chains and static keys are a more performant combination; more study is needed here.

The following pieces are necessary to hook a certain point in the filesystem:

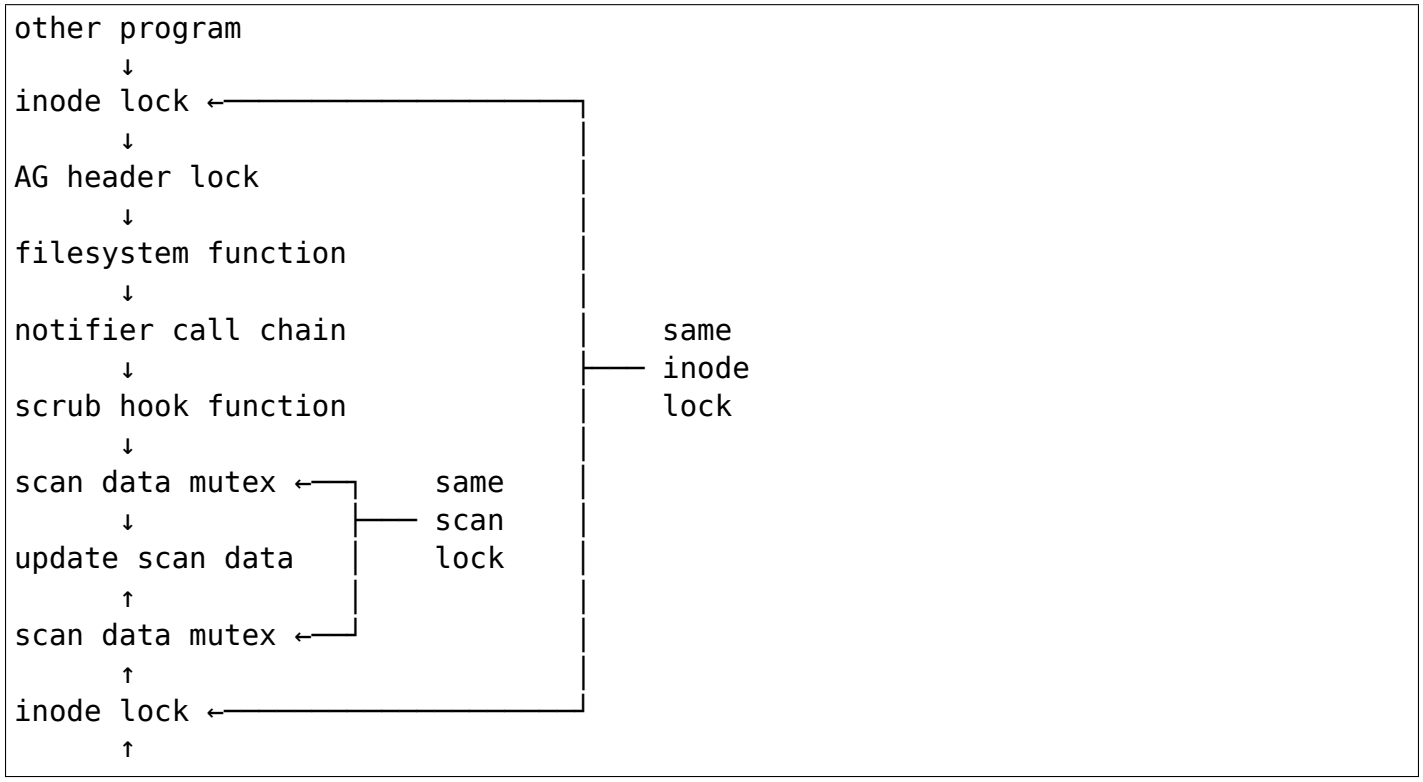
- A `struct xfs_hooks` object must be embedded in a convenient place such as a well-known incore filesystem object.
- Each hook must define an action code and a structure containing more context about the action.

- Hook providers should provide appropriate wrapper functions and structs around the `xfs_hooks` and `xfs_hook` objects to take advantage of type checking to ensure correct usage.
- A callsite in the regular filesystem code must be chosen to call `xfs_hooks_call` with the action code and data structure. This place should be adjacent to (and not earlier than) the place where the filesystem update is committed to the transaction. In general, when the filesystem calls a hook chain, it should be able to handle sleeping and should not be vulnerable to memory reclaim or locking recursion. However, the exact requirements are very dependent on the context of the hook caller and the callee.
- The online fsck function should define a structure to hold scan data, a lock to coordinate access to the scan data, and a `struct xfs_hook` object. The scanner function and the regular filesystem code must acquire resources in the same order; see the next section for details.
- The online fsck code must contain a C function to catch the hook action code and data structure. If the object being updated has already been visited by the scan, then the hook information must be applied to the scan data.
- Prior to unlocking inodes to start the scan, online fsck must call `xfs_hooks_setup` to initialize the `struct xfs_hook`, and `xfs_hooks_add` to enable the hook.
- Online fsck must call `xfs_hooks_del` to disable the hook once the scan is complete.

The number of hooks should be kept to a minimum to reduce complexity. Static keys are used to reduce the overhead of filesystem hooks to nearly zero when online fsck is not running.

### Live Updates During a Scan

The code paths of the online fsck scanning code and the *hooked* filesystem code look like this:



```
scrub function
      ↑
inode scanner
      ↑
xfs_scrub
```

These rules must be followed to ensure correct interactions between the checking code and the code making an update to the filesystem:

- Prior to invoking the notifier call chain, the filesystem function being hooked must acquire the same lock that the scrub scanning function acquires to scan the inode.
- The scanning function and the scrub hook function must coordinate access to the scan data by acquiring a lock on the scan data.
- Scrub hook function must not add the live update information to the scan observations unless the inode being updated has already been scanned. The scan coordinator has a helper predicate (`xchk_iscan_want_live_update`) for this.
- Scrub hook functions must not change the caller's state, including the transaction that it is running. They must not acquire any resources that might conflict with the filesystem function being hooked.
- The hook function can abort the inode scan to avoid breaking the other rules.

The inode scan APIs are pretty simple:

- `xchk_iscan_start` starts a scan
- `xchk_iscan_iter` grabs a reference to the next inode in the scan or returns zero if there is nothing left to scan
- `xchk_iscan_want_live_update` to decide if an inode has already been visited in the scan. This is critical for hook functions to decide if they need to update the in-memory scan information.
- `xchk_iscan_mark_visited` to mark an inode as having been visited in the scan
- `xchk_iscan_teardown` to finish the scan

This functionality is also a part of the inode scanner series.

### Case Study: Quota Counter Checking

It is useful to compare the mount time quotacheck code to the online repair quotacheck code. Mount time quotacheck does not have to contend with concurrent operations, so it does the following:

1. Make sure the ondisk dquots are in good enough shape that all the incore dquots will actually load, and zero the resource usage counters in the ondisk buffer.
2. Walk every inode in the filesystem. Add each file's resource usage to the incore dquot.
3. Walk each incore dquot. If the incore dquot is not being flushed, add the ondisk buffer backing the incore dquot to a delayed write (delwri) list.
4. Write the buffer list to disk.

Like most online fsck functions, online quotacheck can't write to regular filesystem objects until the newly collected metadata reflect all filesystem state. Therefore, online quotacheck records file resource usage to a shadow dquot index implemented with a sparse `xfarray`, and only writes to the real dquots once the scan is complete. Handling transactional updates is tricky because quota resource usage updates are handled in phases to minimize contention on dquots:

1. The inodes involved are joined and locked to a transaction.
2. For each dquot attached to the file:
    a. The dquot is locked.
    b. A quota reservation is added to the dquot's resource usage. The reservation is recorded in the transaction.
    c. The dquot is unlocked.
3. Changes in actual quota usage are tracked in the transaction.
4. At transaction commit time, each dquot is examined again:
    a. The dquot is locked again.
    b. Quota usage changes are logged and unused reservation is given back to the dquot.
    c. The dquot is unlocked.

For online quotacheck, hooks are placed in steps 2 and 4. The step 2 hook creates a shadow version of the transaction dquot context (`dqtrx`) that operates in a similar manner to the regular code. The step 4 hook commits the shadow `dqtrx` changes to the shadow dquots. Notice that both hooks are called with the inode locked, which is how the live update coordinates with the inode scanner.

The quotacheck scan looks like this:

1. Set up a coordinated inode scan.
2. For each inode returned by the inode scan iterator:
    a. Grab and lock the inode.
    b. Determine that inode's resource usage (data blocks, inode counts, realtime blocks) and add that to the shadow dquots for the user, group, and project ids associated with the inode.
    c. Unlock and release the inode.
3. For each dquot in the system:
    a. Grab and lock the dquot.
    b. Check the dquot against the shadow dquots created by the scan and updated by the live hooks.

Live updates are key to being able to walk every quota record without needing to hold any locks for a long duration. If repairs are desired, the real and shadow dquots are locked and their resource counts are set to the values in the shadow dquot.

The proposed patchset is the online quotacheck series.

### Case Study: File Link Count Checking

File link count checking also uses live update hooks. The coordinated inode scanner is used to visit all directories on the filesystem, and per-file link count records are stored in a sparse `xfarray` indexed by inumber. During the scanning phase, each entry in a directory generates observation data as follows:

1. If the entry is a dotdot (`'..'`) entry of the root directory, the directory's parent link count is bumped because the root directory's dotdot entry is self referential.
2. If the entry is a dotdot entry of a subdirectory, the parent's backref count is bumped.
3. If the entry is neither a dot nor a dotdot entry, the target file's parent count is bumped.
4. If the target is a subdirectory, the parent's child link count is bumped.

A crucial point to understand about how the link count inode scanner interacts with the live update hooks is that the scan cursor tracks which *parent* directories have been scanned. In other words, the live updates ignore any update about $\texttt{A} \rightarrow \texttt{B}$ when A has not been scanned, even if B has been scanned. Furthermore, a subdirectory A with a dotdot entry pointing back to B is accounted as a backref counter in the shadow data for A, since child dotdot entries affect the parent's link count. Live update hooks are carefully placed in all parts of the filesystem that create, change, or remove directory entries, since those operations involve bumplink and droplink.

For any file, the correct link count is the number of parents plus the number of child subdirectories. Non-directories never have children of any kind. The backref information is used to detect inconsistencies in the number of links pointing to child subdirectories and the number of dotdot entries pointing back.

After the scan completes, the link count of each file can be checked by locking both the inode and the shadow data, and comparing the link counts. A second coordinated inode scan cursor is used for comparisons. Live updates are key to being able to walk every inode without needing to hold any locks between inodes. If repairs are desired, the inode's link count is set to the value in the shadow information. If no parents are found, the file must be *reparented* to the orphanage to prevent the file from being lost forever.

The proposed patchset is the file link count repair series.

### Case Study: Rebuilding Reverse Mapping Records

Most repair functions follow the same pattern: lock filesystem resources, walk the surviving ondisk metadata looking for replacement metadata records, and use an *in-memory array* to store the gathered observations. The primary advantage of this approach is the simplicity and modularity of the repair code -- code and data are entirely contained within the scrub module, do not require hooks in the main filesystem, and are usually the most efficient in memory use. A secondary advantage of this repair approach is atomicity -- once the kernel decides a structure is corrupt, no other threads can access the metadata until the kernel finishes repairing and revalidating the metadata.

For repairs going on within a shard of the filesystem, these advantages outweigh the delays inherent in locking the shard while repairing parts of the shard. Unfortunately, repairs to the reverse mapping btree cannot use the ”standard” btree repair strategy because it must scan every space mapping of every fork of every file in the filesystem, and the filesystem cannot stop.

Therefore, rmap repair foregoes atomicity between scrub and repair. It combines a *coordinated inode scanner*, *live update hooks*, and an *in-memory rmap btree* to complete the scan for reverse mapping records.

1. Set up an xfbtree to stage rmap records.
2. While holding the locks on the AGI and AGF buffers acquired during the scrub, generate reverse mappings for all AG metadata: inodes, btrees, CoW staging extents, and the internal log.
3. Set up an inode scanner.
4. Hook into rmap updates for the AG being repaired so that the live scan data can receive updates to the rmap btree from the rest of the filesystem during the file scan.
5. For each space mapping found in either fork of each file scanned, decide if the mapping matches the AG of interest. If so:
    a. Create a btree cursor for the in-memory btree.
    b. Use the rmap code to add the record to the in-memory btree.
    c. Use the *special commit function* to write the xfbtree changes to the xfile.
6. For each live update received via the hook, decide if the owner has already been scanned. If so, apply the live update into the scan data:
    a. Create a btree cursor for the in-memory btree.
    b. Replay the operation into the in-memory btree.
    c. Use the *special commit function* to write the xfbtree changes to the xfile. This is performed with an empty transaction to avoid changing the caller's state.
7. When the inode scan finishes, create a new scrub transaction and relock the two AG headers.
8. Compute the new btree geometry using the number of rmap records in the shadow btree, like all other btree rebuilding functions.
9. Allocate the number of blocks computed in the previous step.
10. Perform the usual btree bulk loading and commit to install the new rmap btree.
11. Reap the old rmap btree blocks as discussed in the case study about how to *reap after rmap btree repair*.
12. Free the xfbtree now that it not needed.

The proposed patchset is the rmap repair series.

### Staging Repairs with Temporary Files on Disk

XFS stores a substantial amount of metadata in file forks: directories, extended attributes, symbolic link targets, free space bitmaps and summary information for the realtime volume, and quota records. File forks map 64-bit logical file fork space extents to physical storage space extents, similar to how a memory management unit maps 64-bit virtual addresses to physical memory addresses. Therefore, file-based tree structures (such as directories and extended attributes) use blocks mapped in the file fork offset address space that point to other blocks mapped within that same address space, and file-based linear structures (such as bitmaps and quota records) compute array element offsets in the file fork offset address space.

Because file forks can consume as much space as the entire filesystem, repairs cannot be staged in memory, even when a paging scheme is available. Therefore, online repair of file-based metadata createas a temporary file in the XFS filesystem, writes a new structure at the correct offsets into the temporary file, and atomically swaps the fork mappings (and hence the fork contents) to commit the repair. Once the repair is complete, the old fork can be reaped as necessary; if the system goes down during the reap, the iunlink code will delete the blocks during log recovery.

**Note**: All space usage and inode indices in the filesystem *must* be consistent to use a temporary file safely! This dependency is the reason why online repair can only use pageable kernel memory to stage ondisk space usage information.

Swapping metadata extents with a temporary file requires the owner field of the block headers to match the file being repaired and not the temporary file. The directory, extended attribute, and symbolic link functions were all modified to allow callers to specify owner numbers explicitly.

There is a downside to the reaping process -- if the system crashes during the reap phase and the fork extents are crosslinked, the iunlink processing will fail because freeing space will find the extra reverse mappings and abort.

Temporary files created for repair are similar to `O_TMPFILE` files created by userspace. They are not linked into a directory and the entire file will be reaped when the last reference to the file is lost. The key differences are that these files must have no access permission outside the kernel at all, they must be specially marked to prevent them from being opened by handle, and they must never be linked into the directory tree.

**Historical Sidebar**:

In the initial iteration of file metadata repair, the damaged metadata blocks would be scanned for salvageable data; the extents in the file fork would be reaped; and then a new structure would be built in its place. This strategy did not survive the introduction of the atomic repair requirement expressed earlier in this document.

The second iteration explored building a second structure at a high offset in the fork from the salvage data, reaping the old extents, and using a `COLLAPSE_RANGE` operation to slide the new extents into place.

This had many drawbacks:

- Array structures are linearly addressed, and the regular filesystem codebase does not have the concept of a linear offset that could be applied to the record offset computation to build an alternate copy.
- Extended attributes are allowed to use the entire attr fork offset address space.
- Even if repair could build an alternate copy of a data structure in a different part of the fork address space, the atomic repair commit requirement means that online repair would have to be able to perform a log assisted `COLLAPSE_RANGE` operation to ensure that the old structure was completely replaced.
- A crash after construction of the secondary tree but before the range collapse would leave unreachable blocks in the file fork. This would likely confuse things further.
- Reaping blocks after a repair is not a simple operation, and initiating a reap operation from a restarted range collapse operation during log recovery is daunting.
- Directory entry blocks and quota records record the file fork offset in the header area of each block. An atomic range collapse operation would have to rewrite this part of each block header. Rewriting a single field in block headers is not a huge problem, but it's something to be aware of.
- Each block in a directory or extended attributes btree index contains sibling and child block pointers. Were the atomic commit to use a range collapse operation, each block would have to be rewritten very carefully to preserve the graph structure. Doing this as part of a range collapse means rewriting a large number of blocks repeatedly, which is not conducive to quick repairs.

This lead to the introduction of temporary file staging.

### Using a Temporary File

Online repair code should use the `xrep_tempfile_create` function to create a temporary file inside the filesystem. This allocates an inode, marks the in-core inode private, and attaches it to the scrub context. These files are hidden from userspace, may not be added to the directory tree, and must be kept private.

Temporary files only use two inode locks: the IOLOCK and the ILOCK. The MMAPLOCK is not needed here, because there must not be page faults from userspace for data fork blocks. The usage patterns of these two locks are the same as for any other XFS file -- access to file data are controlled via the IOLOCK, and access to file metadata are controlled via the ILOCK. Locking helpers are provided so that the temporary file and its lock state can be cleaned up by the scrub context. To comply with the nested locking strategy laid out in the *inode locking* section, it is recommended that scrub functions use the xrep_tempfile_ilock*_nowait lock helpers.

Data can be written to a temporary file by two means:

1. `xrep_tempfile_copyin` can be used to set the contents of a regular temporary file from an xfile.

2. The regular directory, symbolic link, and extended attribute functions can be used to write to the temporary file.

Once a good copy of a data file has been constructed in a temporary file, it must be conveyed to the file being repaired, which is the topic of the next section.

The proposed patches are in the repair temporary files series.

### Atomic Extent Swapping

Once repair builds a temporary file with a new data structure written into it, it must commit the new changes into the existing file. It is not possible to swap the inumbers of two files, so instead the new metadata must replace the old. This suggests the need for the ability to swap extents, but the existing extent swapping code used by the file defragmenting tool `xfs_fsr` is not sufficient for online repair because:

a. When the reverse-mapping btree is enabled, the swap code must keep the reverse mapping information up to date with every exchange of mappings. Therefore, it can only exchange one mapping per transaction, and each transaction is independent.

b. Reverse-mapping is critical for the operation of online fsck, so the old defragmentation code (which swapped entire extent forks in a single operation) is not useful here.

c. Defragmentation is assumed to occur between two files with identical contents. For this use case, an incomplete exchange will not result in a user-visible change in file contents, even if the operation is interrupted.

d. Online repair needs to swap the contents of two files that are by definition *not* identical. For directory and xattr repairs, the user-visible contents might be the same, but the contents of individual blocks may be very different.

e. Old blocks in the file may be cross-linked with another structure and must not reappear if the system goes down mid-repair.

These problems are overcome by creating a new deferred operation and a new type of log intent item to track the progress of an operation to exchange two file ranges. The new deferred operation type chains together the same transactions used by the reverse-mapping extent swap code. The new log item records the progress of the exchange to ensure that once an exchange begins, it will always run to completion, even there are interruptions. The new `XFS_SB_FEAT_INCOMPAT_LOG_ATOMIC_SWAP` log-incompatible feature flag in the superblock protects these new log item records from being replayed on old kernels.

The proposed patchset is the atomic extent swap series.

**Sidebar: Using Log-Incompatible Feature Flags**

Starting with XFS v5, the superblock contains a `sb_features_log_incompat` field to indicate that the log contains records that might not readable by all kernels that could mount this filesystem. In short, log incompat features protect the log contents against kernels that will not understand the contents. Unlike the other superblock feature bits, log incompat bits are ephemeral because an empty (clean) log does not need protection. The log cleans itself after its contents have been committed into the filesystem, either as part of an unmount or because the system is otherwise idle. Because upper level code can be working on a transaction at the same time that the log cleans itself, it is necessary for upper level code to communicate to the log when it is going to use a log incompatible feature.

The log coordinates access to incompatible features through the use of one `struct rw_semaphore` for each feature. The log cleaning code tries to take this rwsem in exclusive mode to clear the bit; if the lock attempt fails, the feature bit remains set. Filesystem code signals its intention to use a log incompat feature in a transaction by calling `xlog_use_incompat_feat`, which takes the rwsem in shared mode. The code supporting a log incompat feature should create wrapper functions to obtain the log feature and call `xfs_add_incompat_log_feature` to set the feature bits in the primary superblock. The superblock update is performed transactionally, so the wrapper to obtain log assistance must be called just prior to the creation of the transaction that uses the functionality. For a file operation, this step must happen after taking the IOLOCK and the MMAPLOCK, but before allocating the transaction. When the transaction is complete, the `xlog_drop_incompat_feat` function is called to release the feature. The feature bit will not be cleared from the superblock until the log becomes clean.

Log-assisted extended attribute updates and atomic extent swaps both use log incompat features and provide convenience wrappers around the functionality.

### Mechanics of an Atomic Extent Swap

Swapping entire file forks is a complex task. The goal is to exchange all file fork mappings between two file fork offset ranges. There are likely to be many extent mappings in each fork, and the edges of the mappings aren't necessarily aligned. Furthermore, there may be other updates that need to happen after the swap, such as exchanging file sizes, inode flags, or conversion of fork data to local format. This is roughly the format of the new deferred extent swap work item:

```
struct xfs_swapext_intent {
    /* Inodes participating in the operation. */
    struct xfs_inode    *sxi_ip1;
    struct xfs_inode    *sxi_ip2;

    /* File offset range information. */
    xfs_fileoff_t       sxi_startoff1;
    xfs_fileoff_t       sxi_startoff2;
    xfs_filblks_t       sxi_blockcount;

    /* Set these file sizes after the operation, unless negative. */
    xfs_fsize_t         sxi_isize1;
    xfs_fsize_t         sxi_isize2;
```

```
    /* XFS_SWAP_EXT_* log operation flags */
    uint64_t            sxi_flags;
};
```

The new log intent item contains enough information to track two logical fork offset ranges: `(inode1, startoff1, blockcount)` and `(inode2, startoff2, blockcount)`. Each step of a swap operation exchanges the largest file range mapping possible from one file to the other. After each step in the swap operation, the two startoff fields are incremented and the blockcount field is decremented to reflect the progress made. The flags field captures behavioral parameters such as swapping the attr fork instead of the data fork and other work to be done after the extent swap. The two isize fields are used to swap the file size at the end of the operation if the file data fork is the target of the swap operation.

When the extent swap is initiated, the sequence of operations is as follows:

1. Create a deferred work item for the extent swap. At the start, it should contain the entirety of the file ranges to be swapped.

2. Call `xfs_defer_finish` to process the exchange. This is encapsulated in `xrep_tempswap_contents` for scrub operations. This will log an extent swap intent item to the transaction for the deferred extent swap work item.

3. Until `sxi_blockcount` of the deferred extent swap work item is zero,

    a. Read the block maps of both file ranges starting at `sxi_startoff1` and `sxi_startoff2`, respectively, and compute the longest extent that can be swapped in a single step. This is the minimum of the two `br_blockcount` s in the mappings. Keep advancing through the file forks until at least one of the mappings contains written blocks. Mutual holes, unwritten extents, and extent mappings to the same physical space are not exchanged.

    For the next few steps, this document will refer to the mapping that came from file 1 as ”map1”, and the mapping that came from file 2 as ”map2”.

    b. Create a deferred block mapping update to unmap map1 from file 1.

    c. Create a deferred block mapping update to unmap map2 from file 2.

    d. Create a deferred block mapping update to map map1 into file 2.

    e. Create a deferred block mapping update to map map2 into file 1.

    f. Log the block, quota, and extent count updates for both files.

    g. Extend the ondisk size of either file if necessary.

    h. Log an extent swap done log item for the extent swap intent log item that was read at the start of step 3.

    i. Compute the amount of file range that has just been covered. This quantity is `(map1.br_startoff + map1.br_blockcount - sxi_startoff1)`, because step 3a could have skipped holes.

    j. Increase the starting offsets of `sxi_startoff1` and `sxi_startoff2` by the number of blocks computed in the previous step, and decrease `sxi_blockcount` by the same quantity. This advances the cursor.

    k. Log a new extent swap intent log item reflecting the advanced state of the work item.

l. Return the proper error code (EAGAIN) to the deferred operation manager to inform it that there is more work to be done. The operation manager completes the deferred work in steps 3b-3e before moving back to the start of step 3.

4. Perform any post-processing. This will be discussed in more detail in subsequent sections.

If the filesystem goes down in the middle of an operation, log recovery will find the most recent unfinished extent swap log intent item and restart from there. This is how extent swapping guarantees that an outside observer will either see the old broken structure or the new one, and never a mismash of both.

### Preparation for Extent Swapping

There are a few things that need to be taken care of before initiating an atomic extent swap operation. First, regular files require the page cache to be flushed to disk before the operation begins, and directio writes to be quiesced. Like any filesystem operation, extent swapping must determine the maximum amount of disk space and quota that can be consumed on behalf of both files in the operation, and reserve that quantity of resources to avoid an unrecoverable out of space failure once it starts dirtying metadata. The preparation step scans the ranges of both files to estimate:

- Data device blocks needed to handle the repeated updates to the fork mappings.
- Change in data and realtime block counts for both files.
- Increase in quota usage for both files, if the two files do not share the same set of quota ids.
- The number of extent mappings that will be added to each file.
- Whether or not there are partially written realtime extents. User programs must never be able to access a realtime file extent that maps to different extents on the realtime volume, which could happen if the operation fails to run to completion.

The need for precise estimation increases the run time of the swap operation, but it is very important to maintain correct accounting. The filesystem must not run completely out of free space, nor can the extent swap ever add more extent mappings to a fork than it can support. Regular users are required to abide the quota limits, though metadata repairs may exceed quota to resolve inconsistent metadata elsewhere.

### Special Features for Swapping Metadata File Extents

Extended attributes, symbolic links, and directories can set the fork format to ”local” and treat the fork as a literal area for data storage. Metadata repairs must take extra steps to support these cases:

- If both forks are in local format and the fork areas are large enough, the swap is performed by copying the incore fork contents, logging both forks, and committing. The atomic extent swap mechanism is not necessary, since this can be done with a single transaction.
- If both forks map blocks, then the regular atomic extent swap is used.
- Otherwise, only one fork is in local format. The contents of the local format fork are converted to a block to perform the swap. The conversion to block format must be done in the same transaction that logs the initial extent swap intent log item. The regular atomic

extent swap is used to exchange the mappings. Special flags are set on the swap operation so that the transaction can be rolled one more time to convert the second file's fork back to local format so that the second file will be ready to go as soon as the ILOCK is dropped.

Extended attributes and directories stamp the owning inode into every block, but the buffer verifiers do not actually check the inode number! Although there is no verification, it is still important to maintain referential integrity, so prior to performing the extent swap, online repair builds every block in the new data structure with the owner field of the file being repaired.

After a successful swap operation, the repair operation must reap the old fork blocks by processing each fork mapping through the standard *file extent reaping* mechanism that is done post-repair. If the filesystem should go down during the reap part of the repair, the iunlink processing at the end of recovery will free both the temporary file and whatever blocks were not reaped. However, this iunlink processing omits the cross-link detection of online repair, and is not completely foolproof.

#### Swapping Temporary File Extents

To repair a metadata file, online repair proceeds as follows:

1. Create a temporary repair file.
2. Use the staging data to write out new contents into the temporary repair file. The same fork must be written to as is being repaired.
3. Commit the scrub transaction, since the swap estimation step must be completed before transaction reservations are made.
4. Call `xrep_tempswap_trans_alloc` to allocate a new scrub transaction with the appropriate resource reservations, locks, and fill out a `struct xfs_swapext_req` with the details of the swap operation.
5. Call `xrep_tempswap_contents` to swap the contents.
6. Commit the transaction to complete the repair.

#### Case Study: Repairing the Realtime Summary File

In the ”realtime” section of an XFS filesystem, free space is tracked via a bitmap, similar to Unix FFS. Each bit in the bitmap represents one realtime extent, which is a multiple of the filesystem block size between 4KiB and 1GiB in size. The realtime summary file indexes the number of free extents of a given size to the offset of the block within the realtime free space bitmap where those free extents begin. In other words, the summary file helps the allocator find free extents by length, similar to what the free space by count (cntbt) btree does for the data section.

The summary file itself is a flat file (with no block headers or checksums!) partitioned into `log2(total rt extents)` sections containing enough 32-bit counters to match the number of blocks in the rt bitmap. Each counter records the number of free extents that start in that bitmap block and can satisfy a power-of-two allocation request.

To check the summary file against the bitmap:

1. Take the ILOCK of both the realtime bitmap and summary files.

2. For each free space extent recorded in the bitmap:
    a. Compute the position in the summary file that contains a counter that represents this free extent.
    b. Read the counter from the xfile.
    c. Increment it, and write it back to the xfile.
3. Compare the contents of the xfile against the ondisk file.

To repair the summary file, write the xfile contents into the temporary file and use atomic extent swap to commit the new contents. The temporary file is then reaped.

The proposed patchset is the realtime summary repair series.

### Case Study: Salvaging Extended Attributes

In XFS, extended attributes are implemented as a namespaced name-value store. Values are limited in size to 64KiB, but there is no limit in the number of names. The attribute fork is unpartitioned, which means that the root of the attribute structure is always in logical block zero, but attribute leaf blocks, dabtree index blocks, and remote value blocks are intermixed. Attribute leaf blocks contain variable-sized records that associate user-provided names with the user-provided values. Values larger than a block are allocated separate extents and written there. If the leaf information expands beyond a single block, a directory/attribute btree (`dabtree`) is created to map hashes of attribute names to entries for fast lookup.

Salvaging extended attributes is done as follows:

1. Walk the attr fork mappings of the file being repaired to find the attribute leaf blocks. When one is found,
    a. Walk the attr leaf block to find candidate keys. When one is found,
        1. Check the name for problems, and ignore the name if there are.
        2. Retrieve the value. If that succeeds, add the name and value to the staging xfarray and xfblob.
2. If the memory usage of the xfarray and xfblob exceed a certain amount of memory or there are no more attr fork blocks to examine, unlock the file and add the staged extended attributes to the temporary file.
3. Use atomic extent swapping to exchange the new and old extended attribute structures. The old attribute blocks are now attached to the temporary file.
4. Reap the temporary file.

The proposed patchset is the extended attribute repair series.

### Fixing Directories

Fixing directories is difficult with currently available filesystem features, since directory entries are not redundant. The offline repair tool scans all inodes to find files with nonzero link count, and then it scans all directories to establish parentage of those linked files. Damaged files and directories are zapped, and files with no parent are moved to the `/lost+found` directory. It does not try to salvage anything.

The best that online repair can do at this time is to read directory data blocks and salvage any dirents that look plausible, correct link counts, and move orphans back into the directory tree. The salvage process is discussed in the case study at the end of this section. The *file link count fsck* code takes care of fixing link counts and moving orphans to the `/lost+found` directory.

### Case Study: Salvaging Directories

Unlike extended attributes, directory blocks are all the same size, so salvaging directories is straightforward:

1. Find the parent of the directory. If the dotdot entry is not unreadable, try to confirm that the alleged parent has a child entry pointing back to the directory being repaired. Otherwise, walk the filesystem to find it.
2. Walk the first partition of data fork of the directory to find the directory entry data blocks. When one is found,
    a. Walk the directory data block to find candidate entries. When an entry is found:
        i. Check the name for problems, and ignore the name if there are.
        ii. Retrieve the inumber and grab the inode. If that succeeds, add the name, inode number, and file type to the staging xfarray and xblob.
3. If the memory usage of the xfarray and xfblob exceed a certain amount of memory or there are no more directory data blocks to examine, unlock the directory and add the staged dirents into the temporary directory. Truncate the staging files.
4. Use atomic extent swapping to exchange the new and old directory structures. The old directory blocks are now attached to the temporary file.
5. Reap the temporary file.

**Future Work Question**: Should repair revalidate the dentry cache when rebuilding a directory?

*Answer*: Yes, it should.

In theory it is necessary to scan all dentry cache entries for a directory to ensure that one of the following apply:

1. The cached dentry reflects an ondisk dirent in the new directory.
2. The cached dentry no longer has a corresponding ondisk dirent in the new directory and the dentry can be purged from the cache.
3. The cached dentry no longer has an ondisk dirent but the dentry cannot be purged. This is the problem case.

Unfortunately, the current dentry cache design doesn't provide a means to walk every child dentry of a specific directory, which makes this a hard problem. There is no known solution.

The proposed patchset is the directory repair series.

### Parent Pointers

A parent pointer is a piece of file metadata that enables a user to locate the file's parent directory without having to traverse the directory tree from the root. Without them, reconstruction of directory trees is hindered in much the same way that the historic lack of reverse space mapping information once hindered reconstruction of filesystem space metadata. The parent pointer feature, however, makes total directory reconstruction possible.

XFS parent pointers include the dirent name and location of the entry within the parent directory. In other words, child files use extended attributes to store pointers to parents in the form `(parent_inum, parent_gen, dirent_pos)` → `(dirent_name)`. The directory checking process can be strengthened to ensure that the target of each dirent also contains a parent pointer pointing back to the dirent. Likewise, each parent pointer can be checked by ensuring that the target of each parent pointer is a directory and that it contains a dirent matching the parent pointer. Both online and offline repair can use this strategy.

**Note**: The ondisk format of parent pointers is not yet finalized.

**Historical Sidebar**:

Directory parent pointers were first proposed as an XFS feature more than a decade ago by SGI. Each link from a parent directory to a child file is mirrored with an extended attribute in the child that could be used to identify the parent directory. Unfortunately, this early implementation had major shortcomings and was never merged into Linux XFS:

1. The XFS codebase of the late 2000s did not have the infrastructure to enforce strong referential integrity in the directory tree. It did not guarantee that a change in a forward link would always be followed up with the corresponding change to the reverse links.
2. Referential integrity was not integrated into offline repair. Checking and repairs were performed on mounted filesystems without taking any kernel or inode locks to coordinate access. It is not clear how this actually worked properly.
3. The extended attribute did not record the name of the directory entry in the parent, so the SGI parent pointer implementation cannot be used to reconnect the directory tree.
4. Extended attribute forks only support 65,536 extents, which means that parent pointer attribute creation is likely to fail at some point before the maximum file link count is achieved.

The original parent pointer design was too unstable for something like a file system repair to depend on. Allison Henderson, Chandan Babu, and Catherine Hoang are working on a second implementation that solves all shortcomings of the first. During 2022, Allison introduced log intent items to track physical manipulations of the extended attribute structures. This solves the referential integrity problem by making it possible to commit a dirent update and a parent pointer update in the same transaction. Chandan increased the maximum extent counts of both data and attribute forks, thereby ensuring that the extended attribute structure can grow to handle the maximum hardlink count of any file.

### Case Study: Repairing Directories with Parent Pointers

Directory rebuilding uses a *coordinated inode scan* and a *directory entry live update hook* as follows:

1. Set up a temporary directory for generating the new directory structure, an xfblob for storing entry names, and an xfarray for stashing directory updates.
2. Set up an inode scanner and hook into the directory entry code to receive updates on directory operations.
3. For each parent pointer found in each file scanned, decide if the parent pointer references the directory of interest. If so:
   a. Stash an addname entry for this dirent in the xfarray for later.
   b. When finished scanning that file, flush the stashed updates to the temporary directory.
4. For each live directory update received via the hook, decide if the child has already been scanned. If so:
   a. Stash an addname or removename entry for this dirent update in the xfarray for later. We cannot write directly to the temporary directory because hook functions are not allowed to modify filesystem metadata. Instead, we stash updates in the xfarray and rely on the scanner thread to apply the stashed updates to the temporary directory.
5. When the scan is complete, atomically swap the contents of the temporary directory and the directory being repaired. The temporary directory now contains the damaged directory structure.
6. Reap the temporary directory.
7. Update the dirent position field of parent pointers as necessary. This may require the queuing of a substantial number of xattr log intent items.

The proposed patchset is the parent pointers directory repair series.

**Unresolved Question**: How will repair ensure that the `dirent_pos` fields match in the reconstructed directory?

*Answer*: There are a few ways to solve this problem:

1. The field could be designated advisory, since the other three values are sufficient to find the entry in the parent. However, this makes indexed key lookup impossible while repairs are ongoing.
2. We could allow creating directory entries at specified offsets, which solves the referential integrity problem but runs the risk that dirent creation will fail due to conflicts with the free space in the directory.

   These conflicts could be resolved by appending the directory entry and amending the xattr code to support updating an xattr key and reindexing the dabtree, though this would have to be performed with the parent directory still locked.
3. Same as above, but remove the old parent pointer entry and add a new one atomically.
4. Change the ondisk xattr format to `(parent_inum, name)` $\rightarrow$ `(parent_gen)`, which would provide the attr name uniqueness that we require, without forcing repair code to update the dirent position. Unfortunately, this requires changes to the xattr code to support attr names as long as 263 bytes.

5. Change the ondisk xattr format to `(parent_inum, hash(name)) → (name, parent_gen)`. If the hash is sufficiently resistant to collisions (e.g. sha256) then this should provide the attr name uniqueness that we require. Names shorter than 247 bytes could be stored directly.

Discussion is ongoing under the parent pointers patch deluge.

### Case Study: Repairing Parent Pointers

Online reconstruction of a file's parent pointer information works similarly to directory reconstruction:

1. Set up a temporary file for generating a new extended attribute structure, an *xfblob<xfblob>* for storing parent pointer names, and an xfarray for stashing parent pointer updates.
2. Set up an inode scanner and hook into the directory entry code to receive updates on directory operations.
3. For each directory entry found in each directory scanned, decide if the dirent references the file of interest. If so:
   a. Stash an addpptr entry for this parent pointer in the xfblob and xfarray for later.
   b. When finished scanning the directory, flush the stashed updates to the temporary directory.
4. For each live directory update received via the hook, decide if the parent has already been scanned. If so:
   a. Stash an addpptr or removepptr entry for this dirent update in the xfarray for later. We cannot write parent pointers directly to the temporary file because hook functions are not allowed to modify filesystem metadata. Instead, we stash updates in the xfarray and rely on the scanner thread to apply the stashed parent pointer updates to the temporary file.
5. Copy all non-parent pointer extended attributes to the temporary file.
6. When the scan is complete, atomically swap the attribute fork of the temporary file and the file being repaired. The temporary file now contains the damaged extended attribute structure.
7. Reap the temporary file.

The proposed patchset is the parent pointers repair series.

### Digression: Offline Checking of Parent Pointers

Examining parent pointers in offline repair works differently because corrupt files are erased long before directory tree connectivity checks are performed. Parent pointer checks are therefore a second pass to be added to the existing connectivity checks:

1. After the set of surviving files has been established (i.e. phase 6), walk the surviving directories of each AG in the filesystem. This is already performed as part of the connectivity checks.

2. For each directory entry found, record the name in an xfblob, and store `(child_ag_inum, parent_inum, parent_gen, dirent_pos)` tuples in a per-AG in-memory slab.
3. For each AG in the filesystem,
    a. Sort the per-AG tuples in order of child_ag_inum, parent_inum, and dirent_pos.
    b. For each inode in the AG,
        1. Scan the inode for parent pointers. Record the names in a per-file xfblob, and store `(parent_inum, parent_gen, dirent_pos)` tuples in a per-file slab.
        2. Sort the per-file tuples in order of parent_inum, and dirent_pos.
        3. Position one slab cursor at the start of the inode's records in the per-AG tuple slab. This should be trivial since the per-AG tuples are in child inumber order.
        4. Position a second slab cursor at the start of the per-file tuple slab.
        5. Iterate the two cursors in lockstep, comparing the parent_ino and dirent_pos fields of the records under each cursor.
            a. Tuples in the per-AG list but not the per-file list are missing and need to be written to the inode.
            b. Tuples in the per-file list but not the per-AG list are dangling and need to be removed from the inode.
            c. For tuples in both lists, update the parent_gen and name components of the parent pointer if necessary.
4. Move on to examining link counts, as we do today.

The proposed patchset is the offline parent pointers repair series.

Rebuilding directories from parent pointers in offline repair is very challenging because it currently uses a single-pass scan of the filesystem during phase 3 to decide which files are corrupt enough to be zapped. This scan would have to be converted into a multi-pass scan:

1. The first pass of the scan zaps corrupt inodes, forks, and attributes much as it does now. Corrupt directories are noted but not zapped.
2. The next pass records parent pointers pointing to the directories noted as being corrupt in the first pass. This second pass may have to happen after the phase 4 scan for duplicate blocks, if phase 4 is also capable of zapping directories.
3. The third pass resets corrupt directories to an empty shortform directory. Free space metadata has not been ensured yet, so repair cannot yet use the directory building code in libxfs.
4. At the start of phase 6, space metadata have been rebuilt. Use the parent pointer information recorded during step 2 to reconstruct the dirents and add them to the now-empty directories.

This code has not yet been constructed.

### The Orphanage

Filesystems present files as a directed, and hopefully acyclic, graph. In other words, a tree. The root of the filesystem is a directory, and each entry in a directory points downwards either to more subdirectories or to non-directory files. Unfortunately, a disruption in the directory graph pointers result in a disconnected graph, which makes files impossible to access via regular path resolution.

Without parent pointers, the directory parent pointer online scrub code can detect a dotdot entry pointing to a parent directory that doesn't have a link back to the child directory and the file link count checker can detect a file that isn't pointed to by any directory in the filesystem. If such a file has a positive link count, the file is an orphan.

With parent pointers, directories can be rebuilt by scanning parent pointers and parent pointers can be rebuilt by scanning directories. This should reduce the incidence of files ending up in `/lost+found`.

When orphans are found, they should be reconnected to the directory tree. Offline fsck solves the problem by creating a directory `/lost+found` to serve as an orphanage, and linking orphan files into the orphanage by using the inumber as the name. Reparenting a file to the orphanage does not reset any of its permissions or ACLs.

This process is more involved in the kernel than it is in userspace. The directory and file link count repair setup functions must use the regular VFS mechanisms to create the orphanage directory with all the necessary security attributes and dentry cache entries, just like a regular directory tree modification.

Orphaned files are adopted by the orphanage as follows:

1. Call `xrep_orphanage_try_create` at the start of the scrub setup function to try to ensure that the lost and found directory actually exists. This also attaches the orphanage directory to the scrub context.

2. If the decision is made to reconnect a file, take the IOLOCK of both the orphanage and the file being reattached. The `xrep_orphanage_iolock_two` function follows the inode locking strategy discussed earlier.

3. Call `xrep_orphanage_compute_blkres` and `xrep_orphanage_compute_name` to compute the new name in the orphanage and the block reservation required.

4. Use `xrep_orphanage_adoption_prep` to reserve resources to the repair transaction.

5. Call `xrep_orphanage_adopt` to reparent the orphaned file into the lost and found, and update the kernel dentry cache.

The proposed patches are in the orphanage adoption series.

## 6. Userspace Algorithms and Data Structures

This section discusses the key algorithms and data structures of the userspace program, `xfs_scrub`, that provide the ability to drive metadata checks and repairs in the kernel, verify file data, and look for other potential problems.

### Checking Metadata

Recall the *phases of fsck work* outlined earlier. That structure follows naturally from the data dependencies designed into the filesystem from its beginnings in 1993. In XFS, there are several groups of metadata dependencies:

a. Filesystem summary counts depend on consistency within the inode indices, the allocation group space btrees, and the realtime volume space information.

b. Quota resource counts depend on consistency within the quota file data forks, inode indices, inode records, and the forks of every file on the system.

c. The naming hierarchy depends on consistency within the directory and extended attribute structures. This includes file link counts.

d. Directories, extended attributes, and file data depend on consistency within the file forks that map directory and extended attribute data to physical storage media.

e. The file forks depends on consistency within inode records and the space metadata indices of the allocation groups and the realtime volume. This includes quota and realtime metadata files.

f. Inode records depends on consistency within the inode metadata indices.

g. Realtime space metadata depend on the inode records and data forks of the realtime metadata inodes.

h. The allocation group metadata indices (free space, inodes, reference count, and reverse mapping btrees) depend on consistency within the AG headers and between all the AG metadata btrees.

i. `xfs_scrub` depends on the filesystem being mounted and kernel support for online fsck functionality.

Therefore, a metadata dependency graph is a convenient way to schedule checking operations in the `xfs_scrub` program:

- Phase 1 checks that the provided path maps to an XFS filesystem and detect the kernel's scrubbing abilities, which validates group (i).
- Phase 2 scrubs groups (g) and (h) in parallel using a threaded workqueue.
- Phase 3 scans inodes in parallel. For each inode, groups (f), (e), and (d) are checked, in that order.
- Phase 4 repairs everything in groups (i) through (d) so that phases 5 and 6 may run reliably.
- Phase 5 starts by checking groups (b) and (c) in parallel before moving on to checking names.
- Phase 6 depends on groups (i) through (b) to find file data blocks to verify, to read them, and to report which blocks of which files are affected.

- Phase 7 checks group (a), having validated everything else.

Notice that the data dependencies between groups are enforced by the structure of the program flow.

### Parallel Inode Scans

An XFS filesystem can easily contain hundreds of millions of inodes. Given that XFS targets installations with large high-performance storage, it is desirable to scrub inodes in parallel to minimize runtime, particularly if the program has been invoked manually from a command line. This requires careful scheduling to keep the threads as evenly loaded as possible.

Early iterations of the `xfs_scrub` inode scanner naïvely created a single workqueue and scheduled a single workqueue item per AG. Each workqueue item walked the inode btree (with `XFS_IOC_INUMBERS`) to find inode chunks and then called bulkstat (`XFS_IOC_BULKSTAT`) to gather enough information to construct file handles. The file handle was then passed to a function to generate scrub items for each metadata object of each inode. This simple algorithm leads to thread balancing problems in phase 3 if the filesystem contains one AG with a few large sparse files and the rest of the AGs contain many smaller files. The inode scan dispatch function was not sufficiently granular; it should have been dispatching at the level of individual inodes, or, to constrain memory consumption, inode btree records.

Thanks to Dave Chinner, bounded workqueues in userspace enable `xfs_scrub` to avoid this problem with ease by adding a second workqueue. Just like before, the first workqueue is seeded with one workqueue item per AG, and it uses INUMBERS to find inode btree chunks. The second workqueue, however, is configured with an upper bound on the number of items that can be waiting to be run. Each inode btree chunk found by the first workqueue's workers are queued to the second workqueue, and it is this second workqueue that queries BULKSTAT, creates a file handle, and passes it to a function to generate scrub items for each metadata object of each inode. If the second workqueue is too full, the workqueue add function blocks the first workqueue's workers until the backlog eases. This doesn't completely solve the balancing problem, but reduces it enough to move on to more pressing issues.

The proposed patchsets are the scrub performance tweaks and the inode scan rebalance series.

### Scheduling Repairs

During phase 2, corruptions and inconsistencies reported in any AGI header or inode btree are repaired immediately, because phase 3 relies on proper functioning of the inode indices to find inodes to scan. Failed repairs are rescheduled to phase 4. Problems reported in any other space metadata are deferred to phase 4. Optimization opportunities are always deferred to phase 4, no matter their origin.

During phase 3, corruptions and inconsistencies reported in any part of a file's metadata are repaired immediately if all space metadata were validated during phase 2. Repairs that fail or cannot be repaired immediately are scheduled for phase 4.

In the original design of `xfs_scrub`, it was thought that repairs would be so infrequent that the `struct xfs_scrub_metadata` objects used to communicate with the kernel could also be used as the primary object to schedule repairs. With recent increases in the number of optimizations possible for a given filesystem object, it became much more memory-efficient to track all

eligible repairs for a given filesystem object with a single repair item. Each repair item represents a single lockable object -- AGs, metadata files, individual inodes, or a class of summary information.

Phase 4 is responsible for scheduling a lot of repair work in as quick a manner as is practical. The *data dependencies* outlined earlier still apply, which means that `xfs_scrub` must try to complete the repair work scheduled by phase 2 before trying repair work scheduled by phase 3. The repair process is as follows:

1. Start a round of repair with a workqueue and enough workers to keep the CPUs as busy as the user desires.
    a. For each repair item queued by phase 2,
        i. Ask the kernel to repair everything listed in the repair item for a given filesystem object.
        ii. Make a note if the kernel made any progress in reducing the number of repairs needed for this object.
        iii. If the object no longer requires repairs, revalidate all metadata associated with this object. If the revalidation succeeds, drop the repair item. If not, requeue the item for more repairs.
    b. If any repairs were made, jump back to 1a to retry all the phase 2 items.
    c. For each repair item queued by phase 3,
        i. Ask the kernel to repair everything listed in the repair item for a given filesystem object.
        ii. Make a note if the kernel made any progress in reducing the number of repairs needed for this object.
        iii. If the object no longer requires repairs, revalidate all metadata associated with this object. If the revalidation succeeds, drop the repair item. If not, requeue the item for more repairs.
    d. If any repairs were made, jump back to 1c to retry all the phase 3 items.
2. If step 1 made any repair progress of any kind, jump back to step 1 to start another round of repair.
3. If there are items left to repair, run them all serially one more time. Complain if the repairs were not successful, since this is the last chance to repair anything.

Corruptions and inconsistencies encountered during phases 5 and 7 are repaired immediately. Corrupt file data blocks reported by phase 6 cannot be recovered by the filesystem.

The proposed patchsets are the repair warning improvements, refactoring of the repair data dependency and object tracking, and the repair scheduling improvement series.

### Checking Names for Confusable Unicode Sequences

If `xfs_scrub` succeeds in validating the filesystem metadata by the end of phase 4, it moves on to phase 5, which checks for suspicious looking names in the filesystem. These names consist of the filesystem label, names in directory entries, and the names of extended attributes. Like most Unix filesystems, XFS imposes the sparest of constraints on the contents of a name:

- Slashes and null bytes are not allowed in directory entries.
- Null bytes are not allowed in userspace-visible extended attributes.
- Null bytes are not allowed in the filesystem label.

Directory entries and attribute keys store the length of the name explicitly ondisk, which means that nulls are not name terminators. For this section, the term "naming domain" refers to any place where names are presented together -- all the names in a directory, or all the attributes of a file.

Although the Unix naming constraints are very permissive, the reality of most modern-day Linux systems is that programs work with Unicode character code points to support international languages. These programs typically encode those code points in UTF-8 when interfacing with the C library because the kernel expects null-terminated names. In the common case, therefore, names found in an XFS filesystem are actually UTF-8 encoded Unicode data.

To maximize its expressiveness, the Unicode standard defines separate control points for various characters that render similarly or identically in writing systems around the world. For example, the character "Cyrillic Small Letter A" U+0430 "а" often renders identically to "Latin Small Letter A" U+0061 "a".

The standard also permits characters to be constructed in multiple ways -- either by using a defined code point, or by combining one code point with various combining marks. For example, the character "Angstrom Sign U+212B "Å" can also be expressed as "Latin Capital Letter A" U+0041 "A" followed by "Combining Ring Above" U+030A "◌̊". Both sequences render identically.

Like the standards that preceded it, Unicode also defines various control characters to alter the presentation of text. For example, the character "Right-to-Left Override" U+202E can trick some programs into rendering "moo\xe2\x80\xaegnp.txt" as "mootxt.png". A second category of rendering problems involves whitespace characters. If the character "Zero Width Space" U+200B is encountered in a file name, the name will render identically to a name that does not have the zero width space.

If two names within a naming domain have different byte sequences but render identically, a user may be confused by it. The kernel, in its indifference to upper level encoding schemes, permits this. Most filesystem drivers persist the byte sequence names that are given to them by the VFS.

Techniques for detecting confusable names are explained in great detail in sections 4 and 5 of the Unicode Security Mechanisms document. When `xfs_scrub` detects UTF-8 encoding in use on a system, it uses the Unicode normalization form NFD in conjunction with the confusable name detection component of libicu to identify names with a directory or within a file's extended attributes that could be confused for each other. Names are also checked for control characters, non-rendering characters, and mixing of bidirectional characters. All of these potential issues are reported to the system administrator during phase 5.

### Media Verification of File Data Extents

The system administrator can elect to initiate a media scan of all file data blocks. This scan after validation of all filesystem metadata (except for the summary counters) as phase 6. The scan starts by calling `FS_IOC_GETFSMAP` to scan the filesystem space map to find areas that are allocated to file data fork extents. Gaps between data fork extents that are smaller than 64k are treated as if they were data fork extents to reduce the command setup overhead. When the space map scan accumulates a region larger than 32MB, a media verification request is sent to the disk as a directio read of the raw block device.

If the verification read fails, `xfs_scrub` retries with single-block reads to narrow down the failure to the specific region of the media and recorded. When it has finished issuing verification requests, it again uses the space mapping ioctl to map the recorded media errors back to metadata structures and report what has been lost. For media errors in blocks owned by files, parent pointers can be used to construct file paths from inode numbers for user-friendly reporting.

## 7. Conclusion and Future Work

It is hoped that the reader of this document has followed the designs laid out in this document and now has some familiarity with how XFS performs online rebuilding of its metadata indices, and how filesystem users can interact with that functionality. Although the scope of this work is daunting, it is hoped that this guide will make it easier for code readers to understand what has been built, for whom it has been built, and why. Please feel free to contact the XFS mailing list with questions.

### FIEXCHANGE_RANGE

As discussed earlier, a second frontend to the atomic extent swap mechanism is a new ioctl call that userspace programs can use to commit updates to files atomically. This frontend has been out for review for several years now, though the necessary refinements to online repair and lack of customer demand mean that the proposal has not been pushed very hard.

### Extent Swapping with Regular User Files

As mentioned earlier, XFS has long had the ability to swap extents between files, which is used almost exclusively by `xfs_fsr` to defragment files. The earliest form of this was the fork swap mechanism, where the entire contents of data forks could be exchanged between two files by exchanging the raw bytes in each inode fork's immediate area. When XFS v5 came along with self-describing metadata, this old mechanism grew some log support to continue rewriting the owner fields of BMBT blocks during log recovery. When the reverse mapping btree was later added to XFS, the only way to maintain the consistency of the fork mappings with the reverse mapping index was to develop an iterative mechanism that used deferred bmap and rmap operations to swap mappings one at a time. This mechanism is identical to steps 2-3 from the procedure above except for the new tracking items, because the atomic extent swap mechanism is an iteration of an existing mechanism and not something totally novel. For the narrow case of file defragmentation, the file contents must be identical, so the recovery guarantees are not much of a gain.

Atomic extent swapping is much more flexible than the existing swapext implementations because it can guarantee that the caller never sees a mix of old and new contents even after a

crash, and it can operate on two arbitrary file fork ranges. The extra flexibility enables several new use cases:

- **Atomic commit of file writes**: A userspace process opens a file that it wants to update. Next, it opens a temporary file and calls the file clone operation to reflink the first file's contents into the temporary file. Writes to the original file should instead be written to the temporary file. Finally, the process calls the atomic extent swap system call (`FIEXCHANGE_RANGE`) to exchange the file contents, thereby committing all of the updates to the original file, or none of them.
- **Transactional file updates**: The same mechanism as above, but the caller only wants the commit to occur if the original file's contents have not changed. To make this happen, the calling process snapshots the file modification and change timestamps of the original file before reflinking its data to the temporary file. When the program is ready to commit the changes, it passes the timestamps into the kernel as arguments to the atomic extent swap system call. The kernel only commits the changes if the provided timestamps match the original file.
- **Emulation of atomic block device writes**: Export a block device with a logical sector size matching the filesystem block size to force all writes to be aligned to the filesystem block size. Stage all writes to a temporary file, and when that is complete, call the atomic extent swap system call with a flag to indicate that holes in the temporary file should be ignored. This emulates an atomic device write in software, and can support arbitrary scattered writes.

### Vectorized Scrub

As it turns out, the *refactoring* of repair items mentioned earlier was a catalyst for enabling a vectorized scrub system call. Since 2018, the cost of making a kernel call has increased considerably on some systems to mitigate the effects of speculative execution attacks. This incentivizes program authors to make as few system calls as possible to reduce the number of times an execution path crosses a security boundary.

With vectorized scrub, userspace pushes to the kernel the identity of a filesystem object, a list of scrub types to run against that object, and a simple representation of the data dependencies between the selected scrub types. The kernel executes as much of the caller's plan as it can until it hits a dependency that cannot be satisfied due to a corruption, and tells userspace how much was accomplished. It is hoped that `io_uring` will pick up enough of this functionality that online fsck can use that instead of adding a separate vectored scrub system call to XFS.

The relevant patchsets are the kernel vectorized scrub and userspace vectorized scrub series.

### Quality of Service Targets for Scrub

One serious shortcoming of the online fsck code is that the amount of time that it can spend in the kernel holding resource locks is basically unbounded. Userspace is allowed to send a fatal signal to the process which will cause `xfs_scrub` to exit when it reaches a good stopping point, but there's no way for userspace to provide a time budget to the kernel. Given that the scrub codebase has helpers to detect fatal signals, it shouldn't be too much work to allow userspace to specify a timeout for a scrub/repair operation and abort the operation if it exceeds budget. However, most repair functions have the property that once they begin to touch ondisk

metadata, the operation cannot be cancelled cleanly, after which a QoS timeout is no longer useful.

### Defragmenting Free Space

Over the years, many XFS users have requested the creation of a program to clear a portion of the physical storage underlying a filesystem so that it becomes a contiguous chunk of free space. Call this free space defragmenter `clearspace` for short.

The first piece the `clearspace` program needs is the ability to read the reverse mapping index from userspace. This already exists in the form of the `FS_IOC_GETFSMAP` ioctl. The second piece it needs is a new fallocate mode (`FALLOC_FL_MAP_FREE_SPACE`) that allocates the free space in a region and maps it to a file. Call this file the "space collector" file. The third piece is the ability to force an online repair.

To clear all the metadata out of a portion of physical storage, clearspace uses the new fallocate map-freespace call to map any free space in that region to the space collector file. Next, clearspace finds all metadata blocks in that region by way of `GETFSMAP` and issues forced repair requests on the data structure. This often results in the metadata being rebuilt somewhere that is not being cleared. After each relocation, clearspace calls the "map free space" function again to collect any newly freed space in the region being cleared.

To clear all the file data out of a portion of the physical storage, clearspace uses the FSMAP information to find relevant file data blocks. Having identified a good target, it uses the `FICLONERANGE` call on that part of the file to try to share the physical space with a dummy file. Cloning the extent means that the original owners cannot overwrite the contents; any changes will be written somewhere else via copy-on-write. Clearspace makes its own copy of the frozen extent in an area that is not being cleared, and uses `FIEDEUPRANGE` (or the *atomic extent swap* feature) to change the target file's data extent mapping away from the area being cleared. When all other mappings have been moved, clearspace reflinks the space into the space collector file so that it becomes unavailable.

There are further optimizations that could apply to the above algorithm. To clear a piece of physical storage that has a high sharing factor, it is strongly desirable to retain this sharing factor. In fact, these extents should be moved first to maximize sharing factor after the operation completes. To make this work smoothly, clearspace needs a new ioctl (`FS_IOC_GETREFCOUNTS`) to report reference count information to userspace. With the refcount information exposed, clearspace can quickly find the longest, most shared data extents in the filesystem, and target them first.

**Future Work Question**: How might the filesystem move inode chunks?

*Answer*: To move inode chunks, Dave Chinner constructed a prototype program that creates a new file with the old contents and then locklessly runs around the filesystem updating directory entries. The operation cannot complete if the filesystem goes down. That problem isn't totally insurmountable: create an inode remapping table hidden behind a jump label, and a log item that tracks the kernel walking the filesystem to update directory entries. The trouble is, the kernel can't do anything about open files, since it cannot revoke them.

**Future Work Question**: Can static keys be used to minimize the cost of supporting `revoke()` on XFS files?

*Answer*: Yes. Until the first revocation, the bailout code need not be in the call path at all.

The relevant patchsets are the kernel freespace defrag and userspace freespace defrag series.

#### Shrinking Filesystems

Removing the end of the filesystem ought to be a simple matter of evacuating the data and metadata at the end of the filesystem, and handing the freed space to the shrink code. That requires an evacuation of the space at end of the filesystem, which is a use of free space defragmentation!

## 3.60 ZoneFS - Zone filesystem for Zoned block devices

### 3.60.1 Introduction

zonefs is a very simple file system exposing each zone of a zoned block device as a file. Unlike a regular POSIX-compliant file system with native zoned block device support (e.g. f2fs), zonefs does not hide the sequential write constraint of zoned block devices to the user. Files representing sequential write zones of the device must be written sequentially starting from the end of the file (append only writes).

As such, zonefs is in essence closer to a raw block device access interface than to a full-featured POSIX file system. The goal of zonefs is to simplify the implementation of zoned block device support in applications by replacing raw block device file accesses with a richer file API, avoiding relying on direct block device file ioctls which may be more obscure to developers. One example of this approach is the implementation of LSM (log-structured merge) tree structures (such as used in RocksDB and LevelDB) on zoned block devices by allowing SSTables to be stored in a zone file similarly to a regular file system rather than as a range of sectors of the entire disk. The introduction of the higher level construct ”one file is one zone” can help reducing the amount of changes needed in the application as well as introducing support for different application programming languages.

#### Zoned block devices

Zoned storage devices belong to a class of storage devices with an address space that is divided into zones. A zone is a group of consecutive LBAs and all zones are contiguous (there are no LBA gaps). Zones may have different types.

- Conventional zones: there are no access constraints to LBAs belonging to conventional zones. Any read or write access can be executed, similarly to a regular block device.
- Sequential zones: these zones accept random reads but must be written sequentially. Each sequential zone has a write pointer maintained by the device that keeps track of the mandatory start LBA position of the next write to the device. As a result of this write constraint, LBAs in a sequential zone cannot be overwritten. Sequential zones must first be erased using a special command (zone reset) before rewriting.

Zoned storage devices can be implemented using various recording and media technologies. The most common form of zoned storage today uses the SCSI Zoned Block Commands (ZBC) and Zoned ATA Commands (ZAC) interfaces on Shingled Magnetic Recording (SMR) HDDs.

Solid State Disks (SSD) storage devices can also implement a zoned interface to, for instance, reduce internal write amplification due to garbage collection. The NVMe Zoned NameSpace (ZNS) is a technical proposal of the NVMe standard committee aiming at adding a zoned storage interface to the NVMe protocol.

### 3.60.2 Zonefs Overview

Zonefs exposes the zones of a zoned block device as files. The files representing zones are grouped by zone type, which are themselves represented by sub-directories. This file structure is built entirely using zone information provided by the device and so does not require any complex on-disk metadata structure.

#### On-disk metadata

zonefs on-disk metadata is reduced to an immutable super block which persistently stores a magic number and optional feature flags and values. On mount, zonefs uses blkdev_report_zones() to obtain the device zone configuration and populates the mount point with a static file tree solely based on this information. File sizes come from the device zone type and write pointer position managed by the device itself.

The super block is always written on disk at sector 0. The first zone of the device storing the super block is never exposed as a zone file by zonefs. If the zone containing the super block is a sequential zone, the mkzonefs format tool always ”finishes” the zone, that is, it transitions the zone to a full state to make it read-only, preventing any data write.

#### Zone type sub-directories

Files representing zones of the same type are grouped together under the same sub-directory automatically created on mount.

For conventional zones, the sub-directory ”cnv” is used. This directory is however created if and only if the device has usable conventional zones. If the device only has a single conventional zone at sector 0, the zone will not be exposed as a file as it will be used to store the zonefs super block. For such devices, the ”cnv” sub-directory will not be created.

For sequential write zones, the sub-directory ”seq” is used.

These two directories are the only directories that exist in zonefs. Users cannot create other directories and cannot rename nor delete the ”cnv” and ”seq” sub-directories.

The size of the directories indicated by the st_size field of struct stat, obtained with the stat() or fstat() system calls, indicates the number of files existing under the directory.

#### Zone files

Zone files are named using the number of the zone they represent within the set of zones of a particular type. That is, both the ”cnv” and ”seq” directories contain files named ”0”, ”1”, ”2”, ... The file numbers also represent increasing zone start sector on the device.

All read and write operations to zone files are not allowed beyond the file maximum size, that is, beyond the zone capacity. Any access exceeding the zone capacity is failed with the -EFBIG error.

Creating, deleting, renaming or modifying any attribute of files and sub-directories is not allowed.

The number of blocks of a file as reported by stat() and fstat() indicates the capacity of the zone file, or in other words, the maximum file size.

### Conventional zone files

The size of conventional zone files is fixed to the size of the zone they represent. Conventional zone files cannot be truncated.

These files can be randomly read and written using any type of I/O operation: buffered I/Os, direct I/Os, memory mapped I/Os (mmap), etc. There are no I/O constraint for these files beyond the file size limit mentioned above.

### Sequential zone files

The size of sequential zone files grouped in the ”seq” sub-directory represents the file's zone write pointer position relative to the zone start sector.

Sequential zone files can only be written sequentially, starting from the file end, that is, write operations can only be append writes. Zonefs makes no attempt at accepting random writes and will fail any write request that has a start offset not corresponding to the end of the file, or to the end of the last write issued and still in-flight (for asynchronous I/O operations).

Since dirty page writeback by the page cache does not guarantee a sequential write pattern, zonefs prevents buffered writes and writeable shared mappings on sequential files. Only direct I/O writes are accepted for these files. zonefs relies on the sequential delivery of write I/O requests to the device implemented by the block layer elevator. An elevator implementing the sequential write feature for zoned block device (ELEVATOR_F_ZBD_SEQ_WRITE elevator feature) must be used. This type of elevator (e.g. mq-deadline) is set by default for zoned block devices on device initialization.

There are no restrictions on the type of I/O used for read operations in sequential zone files. Buffered I/Os, direct I/Os and shared read mappings are all accepted.

Truncating sequential zone files is allowed only down to 0, in which case, the zone is reset to rewind the file zone write pointer position to the start of the zone, or up to the zone capacity, in which case the file's zone is transitioned to the FULL state (finish zone operation).

### Format options

Several optional features of zonefs can be enabled at format time.

- Conventional zone aggregation: ranges of contiguous conventional zones can be aggregated into a single larger file instead of the default one file per zone.
- File ownership: The owner UID and GID of zone files is by default 0 (root) but can be changed to any valid UID/GID.
- File access permissions: the default 640 access permissions can be changed.

### IO error handling

Zoned block devices may fail I/O requests for reasons similar to regular block devices, e.g. due to bad sectors. However, in addition to such known I/O failure pattern, the standards governing zoned block devices behavior define additional conditions that result in I/O errors.

- A zone may transition to the read-only condition (BLK_ZONE_COND_READONLY): While the data already written in the zone is still readable, the zone can no longer be written. No user action on the zone (zone management command or read/write access) can change the zone condition back to a normal read/write state. While the reasons for the device to transition a zone to read-only state are not defined by the standards, a typical cause for such transition would be a defective write head on an HDD (all zones under this head are changed to read-only).
- A zone may transition to the offline condition (BLK_ZONE_COND_OFFLINE): An offline zone cannot be read nor written. No user action can transition an offline zone back to an operational good state. Similarly to zone read-only transitions, the reasons for a drive to transition a zone to the offline condition are undefined. A typical cause would be a defective read-write head on an HDD causing all zones on the platter under the broken head to be inaccessible.
- Unaligned write errors: These errors result from the host issuing write requests with a start sector that does not correspond to a zone write pointer position when the write request is executed by the device. Even though zonefs enforces sequential file write for sequential zones, unaligned write errors may still happen in the case of a partial failure of a very large direct I/O operation split into multiple BIOs/requests or asynchronous I/O operations. If one of the write request within the set of sequential write requests issued to the device fails, all write requests queued after it will become unaligned and fail.
- Delayed write errors: similarly to regular block devices, if the device side write cache is enabled, write errors may occur in ranges of previously completed writes when the device write cache is flushed, e.g. on fsync(). Similarly to the previous immediate unaligned write error case, delayed write errors can propagate through a stream of cached sequential data for a zone causing all data to be dropped after the sector that caused the error.

All I/O errors detected by zonefs are notified to the user with an error code return for the system call that triggered or detected the error. The recovery actions taken by zonefs in response to I/O errors depend on the I/O type (read vs write) and on the reason for the error (bad sector, unaligned writes or zone condition change).

- For read I/O errors, zonefs does not execute any particular recovery action, but only if the file zone is still in a good condition and there is no inconsistency between the file inode size and its zone write pointer position. If a problem is detected, I/O error recovery is executed (see below table).
- For write I/O errors, zonefs I/O error recovery is always executed.
- A zone condition change to read-only or offline also always triggers zonefs I/O error recovery.

Zonefs minimal I/O error recovery may change a file size and file access permissions.

- File size changes: Immediate or delayed write errors in a sequential zone file may cause the file inode size to be inconsistent with the amount of data successfully written in the file zone. For instance, the partial failure of a multi-BIO large write operation will cause the zone write pointer to advance partially, even though the entire write operation will be

reported as failed to the user. In such case, the file inode size must be advanced to reflect the zone write pointer change and eventually allow the user to restart writing at the end of the file. A file size may also be reduced to reflect a delayed write error detected on fsync(): in this case, the amount of data effectively written in the zone may be less than originally indicated by the file inode size. After such I/O error, zonefs always fixes the file inode size to reflect the amount of data persistently stored in the file zone.

- Access permission changes: A zone condition change to read-only is indicated with a change in the file access permissions to render the file read-only. This disables changes to the file attributes and data modification. For offline zones, all permissions (read and write) to the file are disabled.

Further action taken by zonefs I/O error recovery can be controlled by the user with the ”errors=xxx” mount option. The table below summarizes the result of zonefs I/O error processing depending on the mount option and on the zone conditions:

| "errors=xxx" mount option | device zone condition | Post error state | | | | |
|---|---|---|---|---|---|---|
| | | file size | file access permissions | | device zone access permissions | |
| | | | read | write | read | write |
| remount-ro (default) | good | fixed | yes | no | yes | yes |
| | read-only | as is | yes | no | yes | no |
| | offline | 0 | no | no | no | no |
| zone-ro | good | fixed | yes | no | yes | yes |
| | read-only | as is | yes | no | yes | no |
| | offline | 0 | no | no | no | no |
| zone-offline | good | 0 | no | no | yes | yes |
| | read-only | 0 | no | no | yes | no |
| | offline | 0 | no | no | no | no |
| repair | good | fixed | yes | yes | yes | yes |
| | read-only | as is | yes | no | yes | no |
| | offline | 0 | no | no | no | no |

Further notes:

- The ”errors=remount-ro” mount option is the default behavior of zonefs I/O error processing if no errors mount option is specified.
- With the ”errors=remount-ro” mount option, the change of the file access permissions to read-only applies to all files. The file system is remounted read-only.
- Access permission and file size changes due to the device transitioning zones to the offline condition are permanent. Remounting or reformatting the device with mkfs.zonefs (mkzonefs) will not change back offline zone files to a good state.
- File access permission changes to read-only due to the device transitioning zones to the read-only condition are permanent. Remounting or reformatting the device will not re-enable file write access.

- File access permission changes implied by the remount-ro, zone-ro and zone-offline mount options are temporary for zones in a good condition. Unmounting and remounting the file system will restore the previous default (format time values) access rights to the files affected.
- The repair mount option triggers only the minimal set of I/O error recovery actions, that is, file size fixes for zones in a good condition. Zones indicated as being read-only or offline by the device still imply changes to the zone file access permissions as noted in the table above.

### Mount options

zonefs defines several mount options: * errors=<behavior> * explicit-open

### ”errors=<behavior>” option

The ”errors=<behavior>” option mount option allows the user to specify zonefs behavior in response to I/O errors, inode size inconsistencies or zone condition changes. The defined behaviors are as follow:

- remount-ro (default)
- zone-ro
- zone-offline
- repair

The run-time I/O error actions defined for each behavior are detailed in the previous section. Mount time I/O errors will cause the mount operation to fail. The handling of read-only zones also differs between mount-time and run-time. If a read-only zone is found at mount time, the zone is always treated in the same manner as offline zones, that is, all accesses are disabled and the zone file size set to 0. This is necessary as the write pointer of read-only zones is defined as invalib by the ZBC and ZAC standards, making it impossible to discover the amount of data that has been written to the zone. In the case of a read-only zone discovered at run-time, as indicated in the previous section. The size of the zone file is left unchanged from its last updated value.

### ”explicit-open” option

A zoned block device (e.g. an NVMe Zoned Namespace device) may have limits on the number of zones that can be active, that is, zones that are in the implicit open, explicit open or closed conditions. This potential limitation translates into a risk for applications to see write IO errors due to this limit being exceeded if the zone of a file is not already active when a write request is issued by the user.

To avoid these potential errors, the ”explicit-open” mount option forces zones to be made active using an open zone command when a file is opened for writing for the first time. If the zone open command succeeds, the application is then guaranteed that write requests can be processed. Conversely, the ”explicit-open” mount option will result in a zone close command being issued to the device on the last close() of a zone file if the zone is not full nor empty.

#### Runtime sysfs attributes

zonefs defines several sysfs attributes for mounted devices. All attributes are user readable and can be found in the directory /sys/fs/zonefs/<dev>/, where <dev> is the name of the mounted zoned block device.

The attributes defined are as follows.

- **max_wro_seq_files**: This attribute reports the maximum number of sequential zone files that can be open for writing. This number corresponds to the maximum number of explicitly or implicitly open zones that the device supports. A value of 0 means that the device has no limit and that any zone (any file) can be open for writing and written at any time, regardless of the state of other zones. When the *explicit-open* mount option is used, zonefs will fail any open() system call requesting to open a sequential zone file for writing when the number of sequential zone files already open for writing has reached the *max_wro_seq_files* limit.
- **nr_wro_seq_files**: This attribute reports the current number of sequential zone files open for writing. When the ”explicit-open” mount option is used, this number can never exceed *max_wro_seq_files*. If the *explicit-open* mount option is not used, the reported number can be greater than *max_wro_seq_files*. In such case, it is the responsibility of the application to not write simultaneously more than *max_wro_seq_files* sequential zone files. Failure to do so can result in write errors.
- **max_active_seq_files**: This attribute reports the maximum number of sequential zone files that are in an active state, that is, sequential zone files that are partially written (not empty nor full) or that have a zone that is explicitly open (which happens only if the *explicit-open* mount option is used). This number is always equal to the maximum number of active zones that the device supports. A value of 0 means that the mounted device has no limit on the number of sequential zone files that can be active.
- **nr_active_seq_files**: This attributes reports the current number of sequential zone files that are active. If *max_active_seq_files* is not 0, then the value of *nr_active_seq_files* can never exceed the value of *nr_active_seq_files*, regardless of the use of the *explicit-open* mount option.

### 3.60.3 Zonefs User Space Tools

The mkzonefs tool is used to format zoned block devices for use with zonefs. This tool is available on Github at:

https://github.com/damien-lemoal/zonefs-tools

zonefs-tools also includes a test suite which can be run against any zoned block device, including null_blk block device created with zoned mode.

### Examples

The following formats a 15TB host-managed SMR HDD with 256 MB zones with the conventional zones aggregation feature enabled:

```
# mkzonefs -o aggr_cnv /dev/sdX
# mount -t zonefs /dev/sdX /mnt
# ls -l /mnt/
total 0
dr-xr-xr-x 2 root root     1 Nov 25 13:23 cnv
dr-xr-xr-x 2 root root 55356 Nov 25 13:23 seq
```

The size of the zone files sub-directories indicate the number of files existing for each type of zones. In this example, there is only one conventional zone file (all conventional zones are aggregated under a single file):

```
# ls -l /mnt/cnv
total 137101312
-rw-r----- 1 root root 140391743488 Nov 25 13:23 0
```

This aggregated conventional zone file can be used as a regular file:

```
# mkfs.ext4 /mnt/cnv/0
# mount -o loop /mnt/cnv/0 /data
```

The ”seq” sub-directory grouping files for sequential write zones has in this example 55356 zones:

```
# ls -lv /mnt/seq
total 14511243264
-rw-r----- 1 root root 0 Nov 25 13:23 0
-rw-r----- 1 root root 0 Nov 25 13:23 1
-rw-r----- 1 root root 0 Nov 25 13:23 2
...
-rw-r----- 1 root root 0 Nov 25 13:23 55354
-rw-r----- 1 root root 0 Nov 25 13:23 55355
```

For sequential write zone files, the file size changes as data is appended at the end of the file, similarly to any regular file system:

```
# dd if=/dev/zero of=/mnt/seq/0 bs=4096 count=1 conv=notrunc oflag=direct
1+0 records in
1+0 records out
4096 bytes (4.1 kB, 4.0 KiB) copied, 0.00044121 s, 9.3 MB/s

# ls -l /mnt/seq/0
-rw-r----- 1 root root 4096 Nov 25 13:23 /mnt/seq/0
```

The written file can be truncated to the zone size, preventing any further write operation:

```
# truncate -s 268435456 /mnt/seq/0
# ls -l /mnt/seq/0
-rw-r----- 1 root root 268435456 Nov 25 13:49 /mnt/seq/0
```

Truncation to 0 size allows freeing the file zone storage space and restart append-writes to the file:

```
# truncate -s 0 /mnt/seq/0
# ls -l /mnt/seq/0
-rw-r----- 1 root root 0 Nov 25 13:49 /mnt/seq/0
```

Since files are statically mapped to zones on the disk, the number of blocks of a file as reported by stat() and fstat() indicates the capacity of the file zone:

```
# stat /mnt/seq/0
File: /mnt/seq/0
Size: 0             Blocks: 524288     IO Block: 4096   regular empty file
Device: 870h/2160d  Inode: 50431       Links: 1
Access: (0640/-rw-r-----)  Uid: (    0/    root)   Gid: (    0/    root)
Access: 2019-11-25 13:23:57.048971997 +0900
Modify: 2019-11-25 13:52:25.553805765 +0900
Change: 2019-11-25 13:52:25.553805765 +0900
Birth: -
```

The number of blocks of the file (”Blocks”) in units of 512B blocks gives the maximum file size of 524288 * 512 B = 256 MB, corresponding to the device zone capacity in this example. Of note is that the ”IO block” field always indicates the minimum I/O size for writes and corresponds to the device physical sector size.

# INDEX

## E



## F

## G



## H



## I



## J

## K



## L



## M



## N



## P

## R



## S

## T



## U



## V



## W



## X