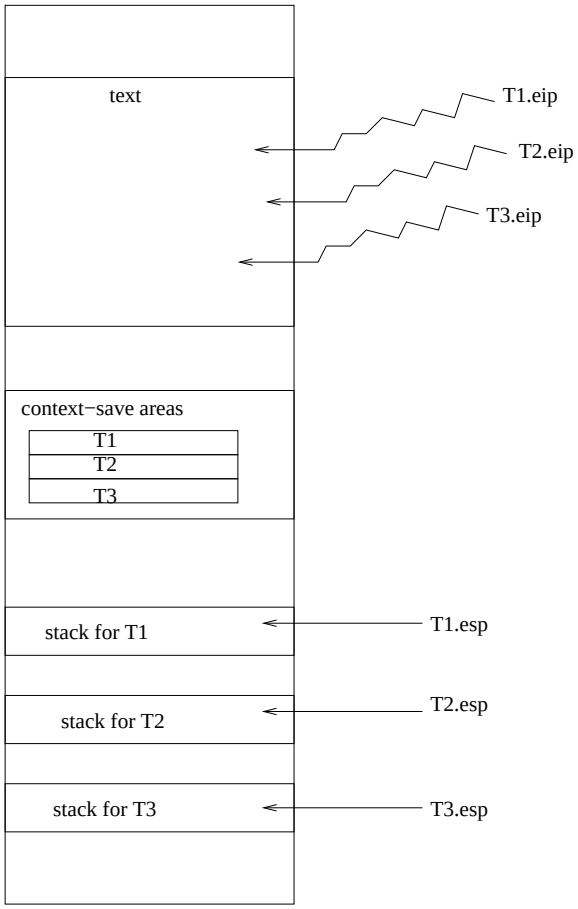
A Task Switch

Consider a multi-threaded user-level program. I.e. one address space, with multiple independent threads of execution within. What defines the **context** of execution on X86-32 architecture?



• The eip register determines the flow of execution (where the next instruction will be fetched).
• The eflags register contains the condition codes and thus affects how conditional branches will happen.
• The esp register controls calling/returning functions because it is where the return address is pushed/popped during CALL / RET. Each thread must have an independent non-overlapping memory area for stack, otherwise there is no meaningful way for functions to exist.
• The ebp register is used in many languages' run-time model to access local variables which live in the stack.
• Other registers such as eax, ebx, etc. are used to store working variables, intermediate results, etc.

So in other words, the values of the CPU registers define the context. Thus from a purely user-level standpoint, assuming that the operating system does not have any underlying support for multi-threaded processes, it would be possible for one task *voluntarily* to give away the sole (virtual) processor to another task by saving all of its registers in some safe place, and then restoring the other tasks' previously-saved register set into the cpu registers. Indeed, this user-level approach was used in very early versions of Linux before thread support was introduced (more than 15 years ago).

Preemptive and Cooperative Multitasking

In previous units, we have defined the terms Pre-emptive and Cooperative (or "Voluntary") multitasking. Since the Linux kernel is a controlled piece of code, it can be trusted to perform cooperative multitasking. However, user-mode programs can't be trusted to relinquish the processor. **Pre-emptive multitasking** means that a currently running task can be forcefully suspended and a context switch made to another, presumably ''better'' task to run.

The trigger for preemption may be based on priority, i.e. a higher-priority task has just become ready to run. It may also be based on time-slicing, in which each task is given a certain amount of time and then another task is given the CPU. This requires a **Periodic Interval Timer** interrupt. Or, it the trigger could be a blend of priority and time. In the Linux kernel, it is the `scheduler` subsystem which makes the decision as to which task to run and when.

In this unit, we shall study the implementation of multi-tasking in the Linux kernel. We have already seen that the kernel is a multi-threaded program of sorts where each thread (task) has its own kernel-mode stack. The situation therefore is analogous to the user-level voluntary task switch described above. The kernel is a *controlled* extension of a process's (task's) context into kernel mode. The kernel code, **acting on behalf of the process**, implements *cooperative* multitasking, which from the standpoint of user processes, appears to be *preemptive* multitasking.

In the Linux kernel, the term **task** is used to mean a schedulable thread of control. When processes are running programs which are multi-threaded, there are multiple tasks running around inside the same process address space. The Linux scheduler works on the basis of tasks, not processes.

The schedule() function

In the Linux kernel, a **task switch** takes place in the `schedule` function, which selects a new task to run (possibly the same task if the system is fairly quiet) and effects the switch

to the new task. Abstractly, `schedule()` is called and then another task has use of the processor. Then at some later time, the original task is scheduled again, and control appears to return transparently from `schedule()`. Internally, `schedule()` maintains a list of tasks which are in a ready to run state (as determined by the `state` field of the `task_struct`), and picks the "best" one. We call this list of ready tasks the **run queue**.

`schedule` can be called in two ways:

• **Directly**: In a synchronous context, `schedule()` is called directly (or through an intermediate function or macro) when the current task wishes to relinquish (voluntarily) the processor because it has reached a blocking state. Examples include a system call which must block waiting for input, or a page fault exception which must block the process until the page fault has been resolved by paging-in from backing store.

• **Indirectly (lazy)**: Whenever control is about to return from an interrupt, fault or system call, the value of the bit flag `TIF_NEED_RESCHED` in the `flags` word of the current `thread_info` structure is examined. We have seen this assembly language code in the previous unit when considering system calls, and similar code exists at the other "return to userland" points. If set, then `schedule()` is invoked. Therefore, when a kernel routine sets `TIF_NEED_RESCHED`, it is requesting that the currently executing task be pre-empted and another task be scheduled. **An asynchronous routine must never call** `schedule` **directly**, because that would leave an unfinished interrupt handler pending. The task switch will occur upon return from the interrupt handler routine. Generally speaking, an interrupt handler routine will set the RESCHED flag if it is waking up a sleeping task which has higher priority than the current task, or when the timer interrupt routine determines that the current task has exhausted its cpu time slice quantum. (The RESCHED flag can also be set in a synchronous context)

Making a context switch -- overview (X86-32)

To make a task (context) switch on a given CPU, the kernel must:

• Save the register context someplace where it won't get clobbered
• Ask the scheduler which task should run next on that CPU
• Adjust scheduler parameters for the new and old task
• Switch the stack pointer to the kernel mode stack for the new task
• Adjust the Task State Segment (TSS) so that the next time control re-enters the kernel from user mode, it does so onto the correct kernel mode stack.
• Adjust the current variable to point to the new task
• If applicable (switching from one user mode process to another vs just one thread to another) switch the user-mode Virtual Address Space.
• And finally, update the instruction pointer (%eip) to "jump" to the previously suspended context point of the "new" task

Kernel Mode Pre-emption

On older Linux kernels, pre-emption could only occur when the task was about to return to user mode from an exception or interrupt handler. **Kernel pre-emption** means that pre-emption can also occur when returning from the interrupt handler back to a kernel control path. Kernel pre-emption can improve performance in real-time applications, because it allows a higher priority task to run immediately. The `preempt_count` per-CPU variable (*it was part of the thread_info struct on older kernels*) controls kernel-mode preemption, and is examined when control is about to return from an interrupt handler back to kernel mode. The kernel macro `preempt_disable` increments this counter. Whenever the counter is non-zero, pre-emption is disabled. `preempt_enable` decrements the counter. Uses of `preempt_disable` can be "nested" and pre-emption is not re-enabled until the outermost kernel routine re-enables. In many places in the kernel, kernel pre-emption is temporarily disabled to prevent leaving things in an inconsistent state.

What schedule() does

An vastly simplified outline of `schedule` is as follows:

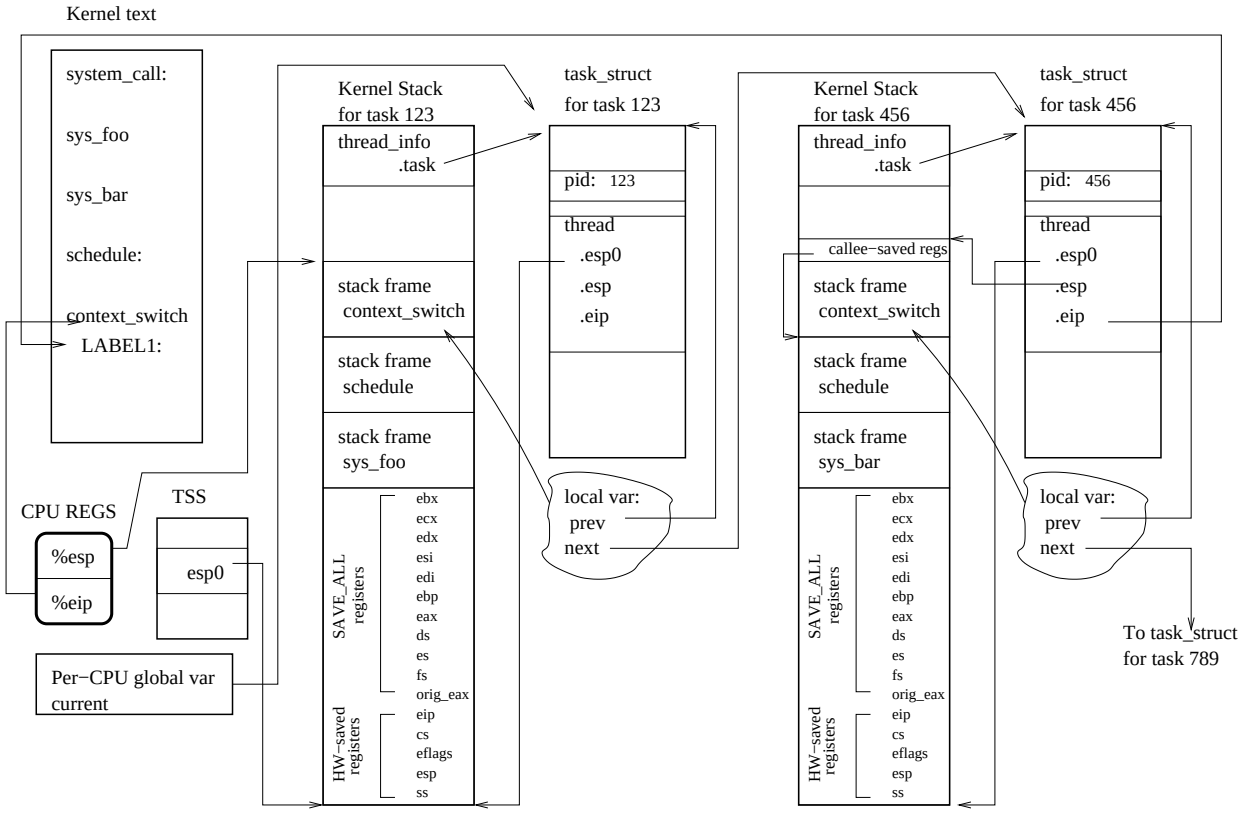
```
                /* from kernel/sched.c*/
schedule()
{
        task_t *prev,*next;
        runqueue_t *rq;

    need_resched:
        preempt_disable(); /* Temporarily disable kernel mode preemption*/
        cpu = smp_processor_id();    /* Get unique id of "this" CPU */
        rq=cpu_rq(cpu);              /* Run queue holding the current task */
        spin_lock(&rq->lock);        /* Lock against others modifying run queue */
        prev=rq->curr;               /* Remember the current task */
        /* ... update a bunch of scheduling statistics ... */
        /* pick_next_task is the "real" scheduler */
        next = pick_next_task(rq);
        prev->thread_info->flags&=~_TIF_NEED_RESCHED;   //Clear flag
        /* ... adjust run time stats of prev ... */
        if (prev!=next)              //Really going to switch now
        {
                sched_info_switch(prev,next);         /* Update timestamps, etc.*/
                rq->curr=next;
                 /* context_switch will also unlock rq*/
                context_switch(rq,prev,next);
                cpu = smp_processor_id();
                rq = cpu_rq(cpu);
        }
        else
                spin_unlock(&rq->lock);
```

```
        preempt_enable();
        if (current_thread_info()->flags & TIF_NEED_RESCHED)
                        goto need_resched;
}

context_switch(runqueue_t *rq, task_t *prev, task_t *next)
{
        prepare_task_switch(rq,prev,next);
        mm = next->mm;
        oldmm = prev->active_mm;
        switch_mm(oldmm,mm,next);             //switch address space

        /* Now it is time for the actual context switch.  This must be
           in assembly language.  The syntax below is simplified */
        asm(
                movl     prev,%eax
                movl     next,%edx
                pushl    %ebx               /* Preserve callee-save register */
                pushl    %edi               /* ebx, edi and esi, plus ebp */
                pushl    %esi               /* on prev's kernel stack */
                pushl    %ebp
 // THREAD is the offset of the .thread member of the task_struct
 // Likewise ESP is the offset of the .esp member of the thread struct
                movl     %esp,(THREAD+ESP)(%eax) /* prev->thread.esp=%esp */
                movl     (THREAD+ESP)(%edx),%esp      /* %esp = next->thread.esp */
 // At this moment we have switched to kernel stack of the NEXT task
                movl     $LABEL1,(THREAD+EIP)(%eax)   /* prev->thread.eip= LABEL1 */
 /* Pushing an address onto the stack and jumping to the function is
        the same as calling the function, but we will return to the
        pushed address instead of the next opcode */
                pushl    (THREAD+EIP)(%edx) /* push next->thread.eip */
                jmp     __switch_to         /* jump to C routine */
 /* When we are scheduled in again, we will come to life at LABEL1.  Yes,
        this is the next opcode.  Why did we bother?  We'll see when we
        look at fork! */
   LABEL1:
                popl     %ebp                        /* restore */
                popl     %esi
                popl     %edi
                popl     %ebx
                movl     %eax,prev          /* return val from __switch_to */
        );

        finish_task_switch(this_rq(),prev);            //Update stats
}


/* The fastcall directive says the arguments are in registers, not stack */
struct task_struct  fastcall *__switch_to(
                                    struct task_struct *prev_p,   //in eax
                                    struct task_struct *next_p)   //in edx
{
                /* Get the thread_structs associated with old and new */
```

```
        struct thread_struct *prev= &prev_p->thread,
                             *next= &next_p->thread;
        int cpu=smp_processor_id();

                /* Get the Task State Segment for this CPU */
        struct tss_struct *tss= &per_cpu(init_tss,cpu);
        tss->x86_tss.esp0 = next->esp0;                 //Set kernel re-entry stack ptr
        /* ... Deal with save/restore of additional registers which
                are normally not touched in kernel mode, such as the
                floating point, mmx/xxm/sse, debug registers, etc.
                These are saved in the ->thread part of the task_struct */
        percpu_write(current_task,next_p);    //update current "variable"
        return prev_p;
}
```

Let us walk through a hypothetical example of making a context switch. Unfortunately, a problem with such examples is the circularity of reasoning. We'll assume that at some time in the past, task #456 had made a system call to (the hypothetical) `sys_bar()`, and that system call reached a blocking state, causing it to call `schedule()`. We will further assume that task #789 was selected by `schedule()` at that time and replaced task #456 on the CPU.

Now, some time has elapsed, and perhaps several intervening tasks have had use of the CPU. Task #123 is currently running in user mode. Somewhere along the way, task #456 was unblocked (e.g. because the input it was waiting for arrived) and thus was placed into the run queue. However, it has still not yet been scheduled. We now come to the beginning point of our example, when task #123 has made a system call `sys_foo()`, and `sys_foo` has to block task #123, thus it has called `schedule()` to yield the CPU. We assume that `schedule()` has selected task #456 as the best task to run next, and has just called `context_switch()`. The situation is depicted as follows:



Examining the kernel stack for task 123, we see the usual `thread_info` structure at the far end of the stack. We see the user's registers which were saved by the hardware and by the `system_call` kernel entrypoint code. We then see a stack frame created for `sys_foo()`, with the return address on the stack being the place where `sys_foo()` was called from the entrypoint code. Likewise, we see that `sys_foo()` called `schedule()`, and that `schedule()` (refer to the code listing above) has local variable `prev` pointing to task #123, and `next` pointing to task #456. Then `schedule()` has called `context_switch()` to effect the actual context switch. The CPU `%esp` stack pointer register is thus pointing to the next location in task 123's kernel stack. The `%eip` program counter register is pointing to the first instruction in `context_switch()`.

Now the following happens within `context_switch()`:

• We need to preserve the general-purpose registers. Recall from the appendix to Unit 7 that on the X86-32 architecture, registers ebx, edi and esi are "callee-saved" meaning that the compiler has generated code which expects that these registers will be safe after a CALL opcode (if the registers are "clobbered" by a called function, that "callee" needs to save and restore them). Since we are about to switch context, those registers would get "clobbered" and would not have the correct values when we are re-scheduled. Therefore it is up to `context_switch` to preserve them. The `%ebp` frame pointer register also needs to be saved. The top of the kernel stack **of task #123** is a safe place to stash these registers because nothing will touch it except task #123. Note that registers eax, ecx and

edx do not need to be saved because these are "caller-saved" registers and are not expected to survive the call to function `context_switch`.

• The current stack pointer (which points to the current top of the kernel stack for task #123, right after we saved the registers above) is saved in `prev->thread.esp` We can't save esp on the stack because how would we find the stack again after we change esp?!

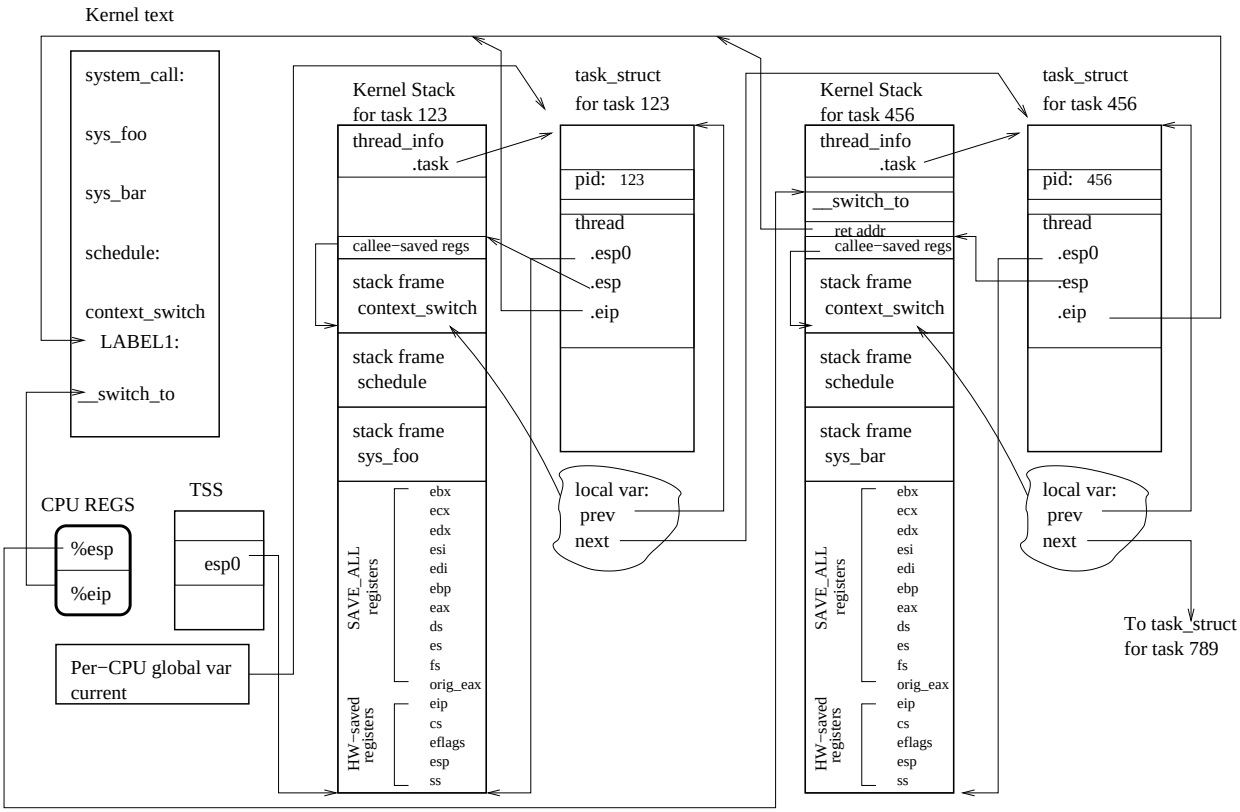
• The address which had been stored in `next->thread.esp` when task #456 was switched-out is also the stack pointer address corresponding to one word beyond the end of the stack frame for `context_switch()`, **BUT** within the kernel stack of task #456. This value is loaded into the `%esp` register, and therefore we are now on the stack of task #456.

• The address of the instruction labeled `LABEL1:` in `context_switch()` is stored in `prev->thread.eip`, i.e. in task #123's task_struct.

• The address which had previously been stored in the `thread.eip` field of task #456's task_struct is, likewise, that same address of the `LABEL1:` label. We'll see an exception to this when dealing with `fork()`. This value is fetched from `thread.eip` and pushed onto the kernel stack `of task #456`.

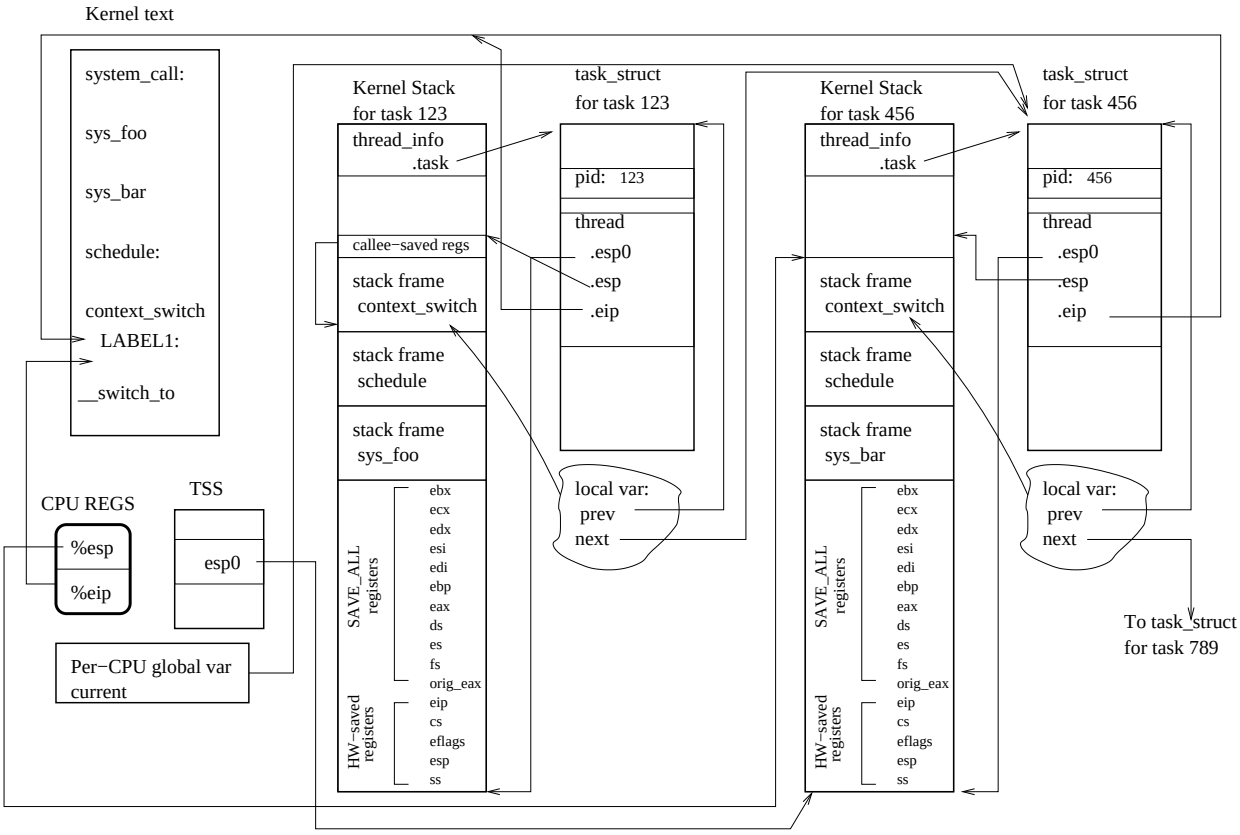
• We jump to the function `__switch_to()`. Normally, C functions are called with a `CALL` instruction, which saves the current `%eip` register on the stack as the return address. Here, the return address is the value which we had loaded from `next->thread.eip`, and manually pushed onto the stack. It so happens in this example that said address corresponds to the next instruction in `context_switch()`, i.e. label `LABEL1:`, but this mechanism allows us to resume execution when switched-in at some other place in the kernel code.

Now let's take a look at things when we have just entered `__switch_to()`:



`__switch_to()` updates the TSS so that at some later time, when task #456 is doing its thing in user mode, and control re-enters the kernel from user mode, it does so with the correct kernel stack pointer, i.e. the kernel stack of task #456.

When `__switch_to()` returns, its return value is the task which we replaced, i.e. the pointer to task struct #123. Thus the local variable `prev` in `context_switch()` is properly set. The `current` variable has likewise been set to point to task #456. We are now ready to return from `context_switch`, and then complete `sys_bar` and return to user mode in task #456:



Process and Thread Creation

Traditionally UNIX processes are self-contained virtual computers with their own private address space and operating system context (e.g. open file descriptors, signal handling). Under Linux, many aspects of the process have been broken out and it is possible to create a new process which selectively shares these attributes with the parent, by using the `clone` system call.

From user-level, the `clone` function is:

```
int clone(int (*fn)(void *), void *child_stack, int flags, void *arg);
```

Unlike `fork`, `clone` begins execution in the child task in the function `fn` which is supplied with the argument `arg`. When this function returns, the child task exits. Before calling `clone`, an area of memory must be allocated and passed as `child_stack`. `clone` as described above is actually a user-level wrapper for the real `clone` system call:

```
int real_clone_syscall(int flags,void *child_stack,...)
```

Note that `fn` is not among these arguments. The real clone system call behaves like fork,

and the user-level wrapper makes sure that the child thread executes `fn(arg)`. The `clone` system call is rarely called directly by a program, and exists primarily to support the threads library (e.g. `pthread_create`). (Because this is an internal system call, the exact argument sequence has varied. What is shown here is a simplification. Programmers should never try to use this raw system call directly!)

Internal to the Linux kernel, there is no distinction between threads, lightweight processes, processes and tasks, and in fact these terms are often used interchangeably in the source code with confusing consequences. Each process has a unique process id. However, the POSIX Threads standard requires that all threads which exist within the same (heavy-weight) process have the same pid as returned by `getpid`. Therefore, Linux introduces the notion of `thread groups`. Each thread has a unique process id (`pid`), and all threads within a heavyweight process have a common `thread group id (tgid)`. It is the `tgid`, not `pid`, which is actually returned by the `getpid` system call. To get the value of the kernel-level `pid`, use the `gettid` system call (get thread id). For conventional single-threaded programs, the `tgid` and `pid` are of course identical.

`flags` contains the signal number to be sent to the parent when the child terminates (usually `SIGCHLD`). This is contained in the low-order byte. Other bitwise flags may be or'd in; an incomplete list follows:
• CLONE_VM: If set, share the virtual address space with the parent. If clear, the child has a copy-on-write private copy of the parent's address space at the instant of clone. If CLONE_VM is set, a new stack area must have been allocated and passed as the `child_stack` parameter, because if not, the child and parent would conflict in their use of the same stack at the same shared virtual address.
• CLONE_SIGHAND: If set, the parent and child will share the signal handling table. Any changes made by one will be reflected in the other's. If not set, the child gets a copy of the signal handler table in effect in the parent at the time of the clone.
• CLONE_FILES: If set, the parent and child will share the open file descriptor table. Therefore, any files open'd in one will be visible using the same fd number in the other. If not set, the child gets a copy of the file descriptor table at the time of the clone.
• CLONE_FS: If set, the parent and child will forever share the following file system information: current root of the filesystem (see `chroot`), current working directory, umask. If not set, the child gets a copy of the parent's information at the time of clone.
• CLONE_PARENT: Affects the notion of who is the parent process after the clone is complete, and thus affects `SIGCHLD` delivery. If set, the parent of the new child is the SAME as the parent of the caller. If not set, the parent of the new child is the caller.
• CLONE_THREAD: If set, the child is put into the same thread group as the caller, and therefore will have the same `tgid`. If not set, the child is placed in a new thread group of which it is the sole member and whose `tgid` is the same as the child's `pid`.

Typically, when `clone` is used to create a new thread with `pthread_create`, all of these CLONE_XX flags are set, so that the new thread exists within the same address

space, and any system calls made by a thread, such as opening a file, affect all of the other threads. Conversely, the `fork` system call sets none of the CLONE_XXX flags, and the new process is thus an independent copy.

The `fork` and `clone` system calls both use the same underlying kernel function `do_fork()`to create a new task/process/thread:

```
/* The usual disclaimer: This code is a highly simplified extract
         of the actual Linux kernel source code */

int sys_fork()
{
         /* plain old fork is the same as clone with no CLONE_XXX flags */
         /* the lowest byte of clone_flags is SIGCHLD, which will be sent */
         /* to the parent when this new process exits.
         /* The child stack pointer is the same as the parent */
         /* since CLONE_VM is not set, the child gets a copy-on-write */
         /* private address space and thus it is OK that the SP are the same */
        return do_fork(SIGCHLD, 0);
}

int sys_clone(unsigned long clone_flags, unsigned long newsp)
{
        return do_fork(clone_flags, newsp);
);
```

Continuing to trace the flow of execution, both clone and fork call the same `do_fork` routine, which in turn calls `copy_process`:

```
long do_fork(unsigned long clone_flags, unsigned long stack_start)
{
        struct task_struct *p;
        int trace = 0;
        long pid = alloc_pidmap();              // Allocate new pid

        if (pid < 0)
                  return -EAGAIN;                       // Oops, out of PIDS
         /* A lot of simplification below */
        p = copy_process(clone_flags, stack_start, pid);
        if (!IS_ERR(p))
                  wake_up_new_task(p,clone_flags);      //place on run queue
         else
         {
                  free_pidmap(pid);
                  pid=PTR_ERR(p);               // Extract errno e.g. ENOMEM
         }
         return pid;
}
```

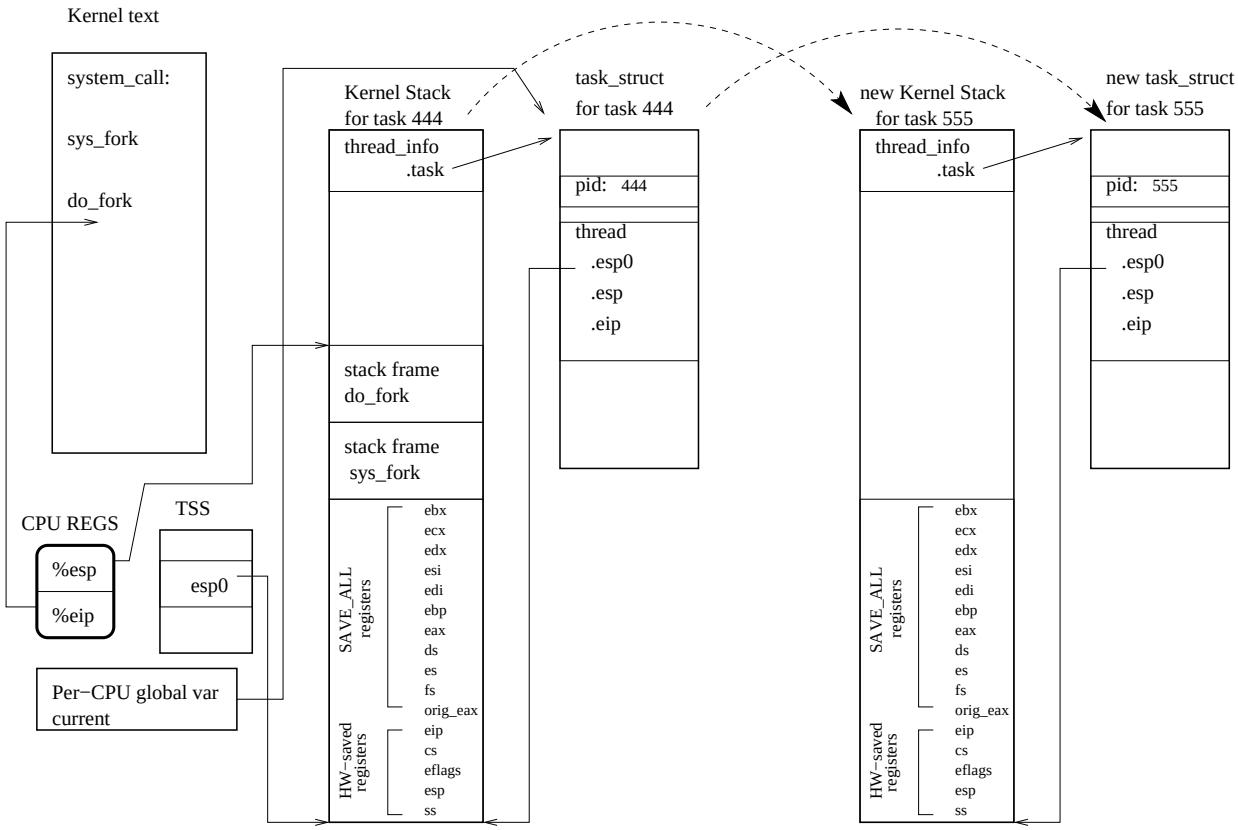
`do_fork` finds an available pid number for the new process or returns `EAGAIN` if there are no free numbers. `copy_process` is invoked, which allocates and returns a pointer to the new `struct task_struct` for the new process. Thanks to the range of valid addresses in kernel memory, a horrendously bad programming practice can be used within the kernel: the returned pointer is overloaded to contain either a valid pointer, or

the (negative) error number, which will thanks to twos-complement be equivalent to a very high (and therefore invalid) pointer address. `copy_process` will return an error code if it did not succeed. Assuming all is well, `wake_up_new_task` is called to place the new task on a run queue so that it may be scheduled for execution.

The real meat is in the `copy_process` function:

```
static struct task_struct *copy_process(unsigned long clone_flags,
                                unsigned long stack_start,
                                int pid)
{
        int retval;
        struct task_struct *p = NULL;

                    /* Some basic sanity checks of the clone flags */
        if ((clone_flags & (CLONE_NEWNS|CLONE_FS)) == (CLONE_NEWNS|CLONE_FS))
                return ERR_PTR(-EINVAL);
        if ((clone_flags & CLONE_THREAD) && !(clone_flags & CLONE_SIGHAND))
                return ERR_PTR(-EINVAL);
        if ((clone_flags & CLONE_SIGHAND) && !(clone_flags & CLONE_VM))
                return ERR_PTR(-EINVAL);
        retval = -ENOMEM;
          /* Make a shallow copy of the task_struct */
          if (!(p=alloc_task_struct())) goto fork_out;
          *p = *current;     /* Shallow structure copy */
          /* Allocate a new kernel stack, set cross-reference pointers */
          struct thread_info *ti;
          if (!(ti=alloc_thread_info())) {
                    free_task_struct(p);
                    p=NULL;
                    goto fork_out;
          }
          p->thread_info=ti;
          ti->task=tsk;
          /* Because the task_structs are the same except for kernel stack addr,
           parent and child now share EVERYTHING.  The copy_XXX routines will
           break the sharing as requested */

          /* ... A bunch of stuff elided here to limit the number of processes
           which a user can spawn, etc. */

          p->pid = pid;
        p->tgid = p->pid;
        if (clone_flags & CLONE_THREAD)
                p->tgid = current->tgid;

        INIT_LIST_HEAD(&p->children);           // Set to empty list
        INIT_LIST_HEAD(&p->sibling);            // Set to empty list
          /* Establish empty signals pending for new task */
        clear_tsk_thread_flag(p, TIF_SIGPENDING);
        init_sigpending(&p->pending);

        p->utime = 0;
```

```
        p->stime = 0;
          acct_clear_integrals(p);     /* zero out per-process accounting
                                fields such as the #system calls, # page faults, etc. */
       /* Perform scheduler related setup. Assign this task to a CPU. */
        sched_fork(p, clone_flags);
        /* now copy all the process information */
          /* each of these copy_xxx functions will look at CLONE_XXX flags to */
          /* figure out if it is a share or a copy.  Some of the options */
          /* as well as the clean up for partial failure have been elided */
        if ((retval = copy_files(clone_flags, p)))       // open files table
                goto fork_out;
        if ((retval = copy_sighand(clone_flags, p)))     // signal handling tbl
                goto fork_out;
        if ((retval = copy_signal(clone_flags, p)))      // pending signals
                goto fork_out;
        if ((retval = copy_mm(clone_flags, p)))          // address space
                goto fork_out;
        copy_thread(clone_flags, stack_start, p);

          /* Mark new task to send signal on exit (normally SIGCHLD), only if
                   this is a new process, as opposed to a thread */
        p->exit_signal = (clone_flags & CLONE_THREAD) ? -1 : (clone_flags & CSIG
NAL);
        p->pdeath_signal = 0;
        p->exit_state = 0;
        if (clone_flags & (CLONE_PARENT|CLONE_THREAD))
                p->real_parent = current->real_parent;
        else
                p->real_parent = current;
        p->parent = p->real_parent;
                   /* Check to see if any signals came in during fork */
          recalc_signal_pending();
          /* ... Insert this process into the pid, pgid, tgid, etc. lists */
          retval=0;          /* We have succeeded! */
          total_forks++;              /* How many forks in life of system */
          nr_threads++;               /* How many current threads */
 fork_out:
          if (retval) return ERR_PTR(retval);
          return p;
}
```

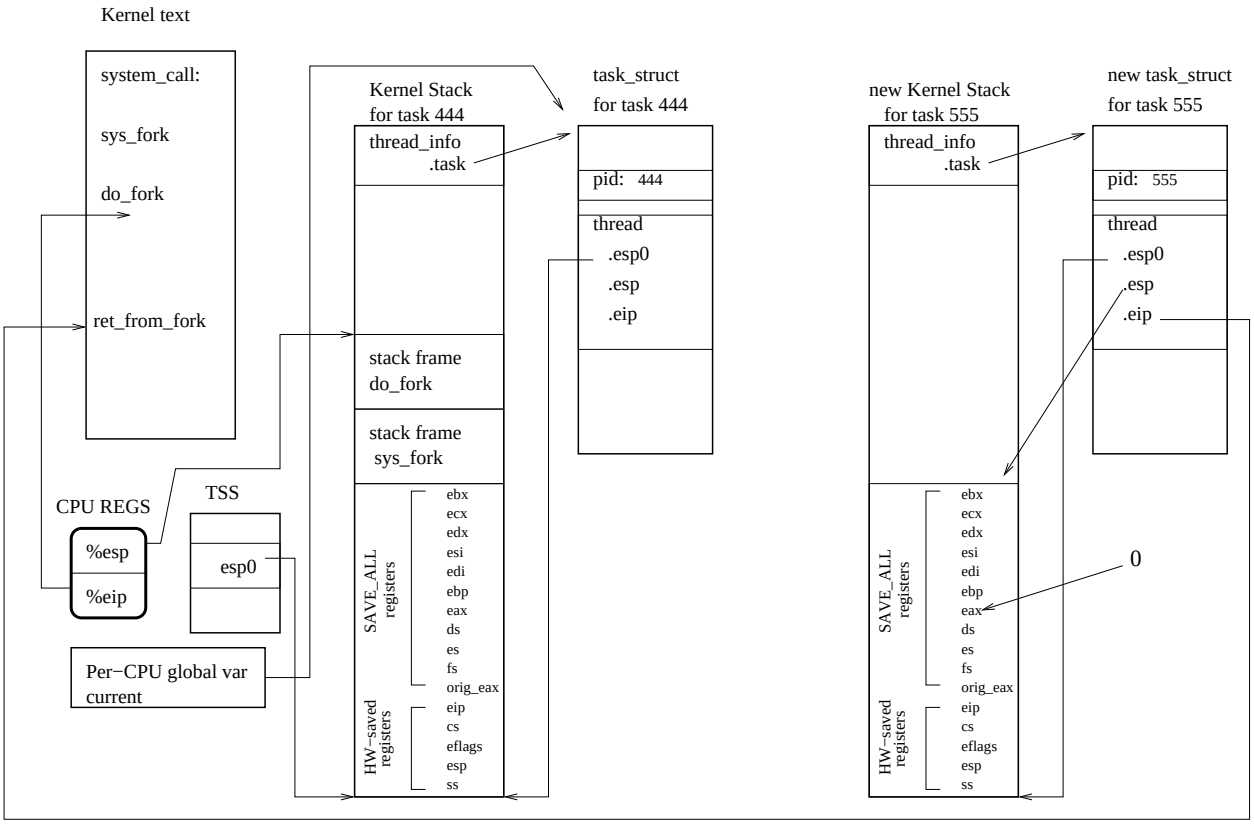


`copy_thread` initializes the brand-new kernel mode stack and task struct, which were created by `dup_task_struct`. The kernel stack for the child is initialized *as if* the child is in the process of making a system call. The registers which were stacked when the parent made the fork/clone system call are copied over, but a 0 is poked into the EAX register slot so that the child will see a 0 return value from the system call.

```
int copy_thread(unsigned long clone_flags, unsigned long stack_start,
        struct task_struct * child)
{
        struct pt_regs *childregs,*parentregs;
        struct task_struct *tsk;
        int err;

          /* Manipulate the user-mode register save area in the kstack */
        childregs = task_pt_regs(child);        // compute reg area address
          parentregs= task_pt_regs(current);
_info)) - 1;
          /* Copy all of the parent's registers to the child's kstack */
        *childregs = *parentregs;
                  /* Set the return value from fork/clone syscall in child to 0*/
        childregs->eax = 0;
                  /* Set the user mode stack pointer for when child first runs */
                  /* For clone, this will be a new value.  For fork, the
                           bulk register copy above set this to the parent's sp */
          if (esp)
```

```
                childregs->esp = esp;

          /* Set the kernel mode stack pointer and stack base pointer */
          /* for when child is first scheduled */
        child->thread.esp = (unsigned long) childregs;
        child->thread.esp0 = (unsigned long) (childregs+1);
                /* When child is scheduled, execution will begin */
                /* in the return_from_fork assembly language function */
        child->thread.eip = (unsigned long) ret_from_fork;
          return 0;
}
```

Once `copy_process` and `do_fork` have completed their work, the child process is ready to run and is waiting on a run queue for its turn to get scheduled. We have already seen how an existing process gets scheduled in with `context_switch`. `do_fork` and its helper functions have now created a new process which is a copy of the current (parent) process (plus or minus the options specified with the CLONE flags) at the exact moment of returning from `fork` or `clone`. The situation just before the parent returns from `do_fork()` is:



Now, when the child process eventually gets scheduled, `context_switch` will find `ret_from_fork` as the new program counter location, rather than the usual address of a local label "LABEL1" within `context_switch`. This is required since the child is not being re-scheduled after having been switched-out, but is being scheduled for the first

time. The necessary stack frame for being in `context_switch()` does not exist in the child. Indeed, the only stack frame is the set of usual registers saved on entry to a system call. Therefore, the child comes to life at the following assembly language code:

```
ret_from_fork:
        pushl    %eax                          #contains prev task_struct ptr
        call     schedule_tail                 #schedule_tail(prev)
        GET_THREAD_INFO(%ebp)                  #mask esp to get thread_info to ebp
        popl %eax
        jmp syscall_exit            #resume exit from system call
```

`schedule_tail()` performs some cleanup actions related to the scheduler. Execution then continues at `syscall_exit` (see previous unit for listing), with the registers and stack looking like an ordinary return from system call. Control returns to user mode, with the return value of the system call being 0 (because the EAX slot in the user-mode saved registers on the child's kernel stack was poked with 0)

Multiprocessor Considerations

On a multi-processor system, each processor is scheduled individually. There can be multiple kernel control paths active at the same time, meaning increased locking complexity. The `current` global variable is maintained as a per-CPU variable. E.g. on a 4 processor system, there are 4 "current" tasks.

Each CPU has an associated run queue. Tasks are placed on a specific run queue when they are created or when they transition from a blocked to a ready state. The scheduling algorithm attempts to balance the loads between CPUs, trading off between the performance hit of unbalanced CPU loads vs the probable cache misses when a task migrates between CPUs.

If there are no runnable tasks for a CPU, it executes the **idle** task. This is either an endless loop waiting for something else to do, or an invocation of power-save mode which is un-done when a hardware interrupt arrives (including possibly an IPI to alert the idle CPU that a new task has been added to its run queue)

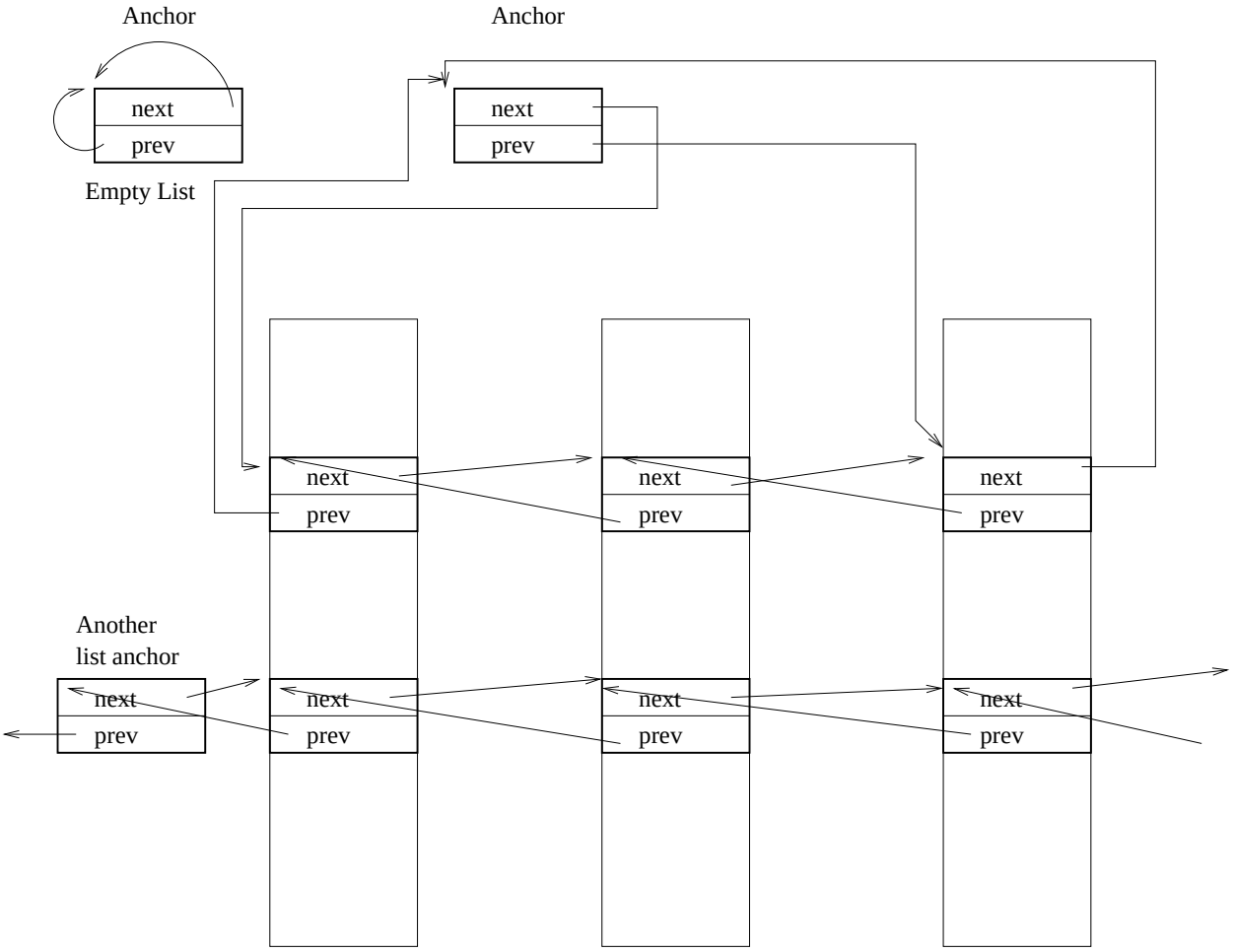
Sidebar: linked lists

Within the Linux kernel, a frequently seen data structure is the `list_head`. The `struct list_head` contains two pointers `next` and `prev` to implement the doubly-linked list. However, rather than pointing at the element itself, these pointers point to the `struct list_head` within each element. The result is that generic list manipulation routines can be used for any list of type FOO, because the offset of the next and prev pointers within an element of FOO is not needed. The macro `list_entry` gets back to the element of type `type` in which `ptr` is the address of the list_head field named `member` within the entry.

```
#define list_entry(ptr, type, member) (                         \
        (type *)( (char *)ptr - (size_t) &((type *)0)->member))
```

Furthermore, the doubly-linked list is circular, meaning there are never NULL pointers. This is advantageous in that it removes a lot of conditional code. The entire list is also represented by a `struct list head`, also known as the list *anchor*. If we call this anchor A, then `A.next` points to the first item on the list, and `A.prev` is the last. It is therefore just as easy to insert an item at either end of the list, which is again used to advantage in many places. The anchor itself is not considered an element of the list. Therefore an empty list is represented by having an anchor where both `next` and `prev` point back to the anchor.

It is frequently the case that a given `struct FOO` inside the kernel contains more than one `list_head` member. Each threads a different circular, doubly-linked list through FOO. E.g. the `task_struct` is on a list of all tasks in the system, and also a distinct list of all children of a particular task, etc.



Sleep and wakeup

Frequently within the kernel a situation is encountered in which a task, executing in a synchronous context, must wait for an event to occur. Examples include:
• System call which needs to wait for an I/O operation to complete.
• System calls involving reading from an empty pipe, or writing to a full pipe.
• Reads from a socket with no data pending, or write to a socket with the buffer full.
• System calls such as `wait` which wait until the state of another task changes.
• "Major" page faults which block until a disk I/O operation completes.

Sleeping on an event and waking up when it occurs involves two context switches, one voluntary, the other potentially pre-emptive. The first, voluntary switch comes when say task "A" encounters the blocking condition in a synchronous context. It places itself on a wait queue, as described below, and then calls `schedule()` voluntarily to yield.

At some later time, the event is satisfied, either in an asynchronous context (e.g. I/O complete interrupt) or a synchronous context (e.g. another process writes to a pipeline). This causes task "A" to be marked as ready to run. A context switch may be forced if the task running at that time, say task "Y", has inferior priority to task "A". Or, a context switch may happen when task "Y" uses up its time slice (see unit 10). In these two cases, the context switch is pre-emptive, and occurs via the `TIF_NEED_RESCHED` flag when task "Y" returns to user mode. Or task A could be scheduled if task Y enters the kernel and blocks.

Wait Queues

For each waitable event, a kernel data structure known as a **wait queue** is defined. It is implemented as a circular, doubly-linked list with some unusual tricks which improve efficiency. This list chains together all of the tasks which are waiting on a specific event.

The wait queue is anchored at a wait queue head structure:

```
typedef struct __wait_queue_head {
         spinlock_t  lock;                    /* mutex */
         struct list_head task_list;          /* anchor */
} wait_queue_head_t;
```

Each element of the list is:

```
typedef struct __wait_queue {
         unsigned int flags;
         struct task_struct *task;
         int (*func)();                                /* callback */
         struct list_head task_list;          /* chains */
} wait_queue_t;
```

A wait queue of type `wait_queue_head_t` contains a circular, doubly-linked list of `wait_queue_t` entries, each of which represents one task waiting on that event.

Kernel code which places tasks on a wait queue is executed in a synchronous situation, because it is running in the context of a process which is being blocked. On a multiprocessor system, multiple kernel routines could be executing in true parallel and potentially trying to insert into the same wait queue (e.g. picture several processes simultaneously blocking on a read from the same input source.) Since this operation is not inherently atomic the `spin_lock` element of the list anchor must be used to protect it.

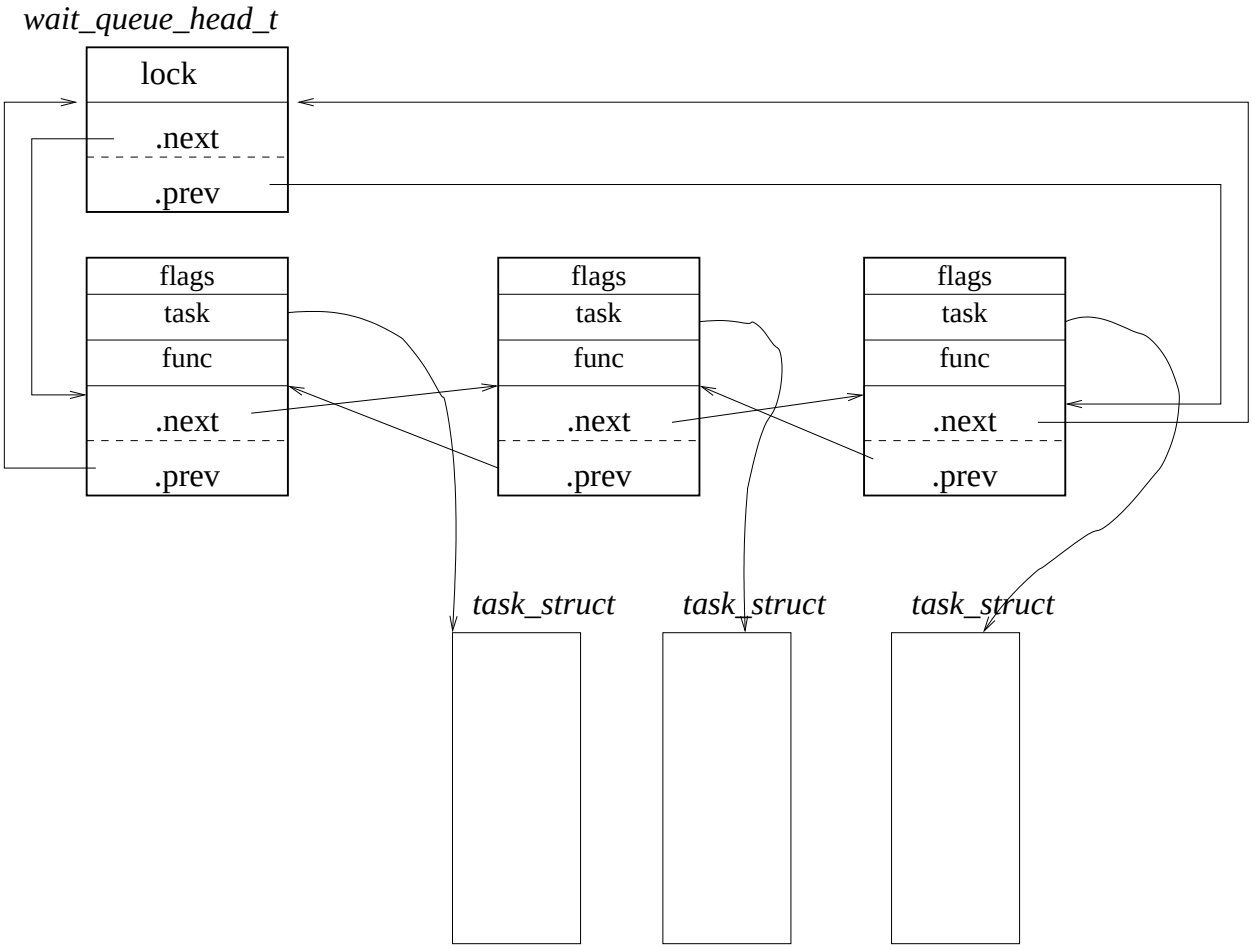
One might ask "why is a spin lock being used to protect the list head, rather than a blocking mutex or semaphore?" The reason is (A) the time that kernel routines are in the critical region is brief and bounded....they need only to obtain the spin lock, manipulate a few pointers to insert a wait queue entry, then release it. (B) Furthermore, code which wakes up tasks is often called from an asynchronous context, e.g. after an I/O completion interrupt. Therefore it can not perform a blocking operation, because it would be blocking a process which is in effect an innocent bystander and which has nothing to do with the wait queue in question.

When a kernel task (in a synchronous situation) is about to perform an operation which could potentially be blocking, it prepares for this by creating a `wait_queue_t` entry (it is acceptable to make this a local variable...the kernel stack is a pretty safe place and control isn't about to leave the function in question) then inserting that entry into the wait queue head, then calling `schedule()` to relinquish the processor. There are two possibilities:

• Exclusive wait: The WQ_FLAG_EXCLUSIVE bit flag will be set in the `flags` field of the `wait_queue_t` entry. At most one exclusive process will be woken up at a time.
• Non-exclusive wait: WQ_FLAG_EXCLUSIVE is clear. All non-exclusive processes on the particular wait queue will be woken up.

Exclusive vs non-exclusive waiting

Exclusive waiters are normally inserted at the tail of the wait queue and non-exclusive waiters at the head. Therefore, if the queue should contain both types of waiters, all of the non-exclusive waiters are awoken first, then at most one exclusive waiter. This situation is actually not common: The choice of exclusive vs non-exclusive wait is made by the waiter at the time it goes to sleep based on whether it makes sense to wake up multiple waiters when the event arrives. If it is likely that only one task could proceed (e.g. waiting on a mutex lock) then the wait will be exclusive to avoid the **"thundering herd"** inefficiency (all tasks wake up, get scheduled, take up CPU time realizing that they have to go back to sleep again). If multiple tasks might be able to proceed (e.g. waiting for a pipe to drain to write more data) then non-exclusive waits make sense.



Wait interrupted by signal

In addition to exclusive vs. non-exclusive waiting, we can distinguish between interruptible and non-interruptible waits. It is useful to allow a signal to interrupt certain potentially long waits (e.g. waiting for a network message to arrive, waiting for a keyboard character). As we shall see, there is cleanup involved when this happens, so the majority of waits within the kernel are non-interruptible. A non-interruptible wait *should* be quick, but sometimes the event takes longer than expected. E.g. most disk I/O operations are non-interruptible. If the disk operation hangs, either because the media has been removed, or there is an I/O error that is being retried, this is manifested as a "freeze" of the process that can not be killed, even with SIGKILL! We hope that the disk driver layers are coded such that eventually the operation will time out and return a "failure" event instead of a "complete" event, which will also wake up the sleeper.

When the task puts itself on a wait queue, it sets its `state` to `TASK_INTERRUPTIBLE` or `TASK_UNINTERRUPTIBLE` to indicate that it is sleeping. Either state prevents the scheduler from placing the task on a run queue, until it is woken up. The

TASK_INTERRUPTIBLE state allows the arrival of a signal to wake up the task too.

Wait queue example

```
// From /usr/src/linux/include/linux/list.h
// self-circular initialization
#define LIST_HEAD_INIT(name) { &(name), &(name) }


// From /usr/src/linux/include/linux/wait.h
// Following uses C-99 structure initialization syntax
#define DEFINE_WAIT(name) \\
         wait_queue_t name = {\\
                  .task=current,\\
                  .func=autoremove_wake_function,\\
                  .task_list=LIST_HEAD_INIT((name).task_list)\\
         }

// The following macro is defined in /usr/src/linux/include/linux/wait.h
// the do { }while(0) construct is a preprocessor trick to make this
// work syntactically as if it were a function
#define __wait_event(wq, condition)                                     \
do {                                                                    \
        DEFINE_WAIT(__wait);                                            \
                                                                        \
        for (;;) {                                                      \
                prepare_to_wait(&wq, &__wait, TASK_UNINTERRUPTIBLE);    \
                if (condition)                                          \
                        break;                                          \
                schedule();                                             \
        }                                                               \
        finish_wait(&wq, &__wait);                                      \
} while (0)


// The following code is abstracted from /usr/src/linux/kernel/wait.c
// there is a similar routine for exclusive waiting
prepare_to_wait( wait_queue_head_t *q, wait_queue_t *wait, int state)
{
         wait->flags&=~WQ_FLAG_EXCLUSIVE;
         spin_lock(&q->lock);                           /* simplified */
         /* This might not be the first time through the loop above, */
         /* in which case the wait_queue entry is already enqueued */
         if (list_empty(&wait->task_list))     /* skip if already enqueued*/
         {
                  list_add(&wait->task_list,&q->task_list);
         }
         // State will be either TASK_INTERRUPTIBLE or TASK_UNINTERRUPTIBLE
         // Either one keeps  this task off the run queue
         current->state=state;
         spin_unlock(&q->lock);
```

```
}

finish_wait(wait_queue_head_t *q, wait_queue_t *wait)
{
          __set_current_state(TASK_RUNNING);    // Basically redundant
          // The next passage basically takes the wait queue entry
          // "wait" and unlinks it from the CLL.  On a multiprocessor
          // system, another processor may also be in the process of
          // waking up this same wait queue entry.
          // list_empty_careful checks that next points to itself
          // AND prev points to itself.
          if (!list_empty_careful(&wait->task_list)) {
                    spin_lock(&q->lock);
                    list_del_init((&wait->task_list);     //remove from list
                    spin_unlock(&q->lock);
          }
}
```

This one of the typical methods as seen within the Linux kernel, however in other cases wait queues are manipulated directly by kernel routines. Unfortunately the Linux kernel is fairly sloppy and inconsistent with synchronization primitives.

Comparing this example with *condition variables* (Unit 6), the use of the mutex to guard against the lost wakeup problem is approached differently. Here in the Linux kernel, the calling task always places itself on the wait queue, before it has tested the condition, THEN it marks itself as being in a sleeping state (TASK_INTERRUPTIBLE or TASK_UNINTERRUPTIBLE in Linux kernel speak). Therefore it can not go to sleep without being on the wait queue. When we look at the wakeup routine below, we'll see that it examines all of the tasks on the given wait queue under the protection of the wait queue's mutex. Therefore no other task can be in the process of inserting itself into the wait queue while the wakeup is happening.

If the waker pulls the wait queue entry and wakes up the sleeper between the time that the sleeper called `prepare_to_wait` and when it called `schedule()`, there will be no missed wakeup. The wait queue lock is not released by the sleeper until it has changed its task state. A waker would gain the lock thereafter and set the task state back to TASK_RUNNING. Thus when the sleeper calls `schedule()` is is already awake and nothing happens.

Waking Up

Once a process puts itself to sleep, it can not be scheduled again until it is woken up. Therefore it is always another task (or an interrupt handler) which wakes up the sleeping process.

```
/*From /usr/src/linux/kernel/sched.c */
/* and /usr/src/linux/kernel/wait.c */
```

```
/* When this is called, the mutex spin lock q->lock has already been grabbed */

static void __wake_up_common(wait_queue_head_t *q, unsigned int mode,
                             int nr_exclusive, int sync, void *key)
{
        struct list_head *tmp, *next;

          // Iterate over wait queue CLL with variable tmp
        list_for_each_safe(tmp, next, &q->task_list) {
                wait_queue_t *curr;
                unsigned flags;
                   // Get back to start of wait queue entry
                curr = list_entry(tmp, wait_queue_t, task_list);
                flags = curr->flags;
                if (curr->func(curr, mode, sync, key) &&
                    (flags & WQ_FLAG_EXCLUSIVE) &&
                    !--nr_exclusive)
                        break;
        }
}

autoremove_wake_function(wait_queue *t, unsigned mode, int sync, void *key)
{
          try_to_wake_up(t->task,mode, wake_flags);
          list_del_init(&wait->task_list);
}

try_to_wake_up(struct task_struct *p,unsigned int state, int wake_flags)
{
          p->state=TASK_WAKING;
          rq=orig_rq=task_rq_lock(p,&flags);     // Get original run queue
          //Execute scheduler class specific hook
          p->sched_class->task_waking(rq,p);
          //Potentially move to different CPU
        cpu = select_task_rq(p, SD_BALANCE_WAKE, wake_flags);
        if (cpu != orig_cpu)
                set_task_cpu(p, cpu);
        rq = __task_rq_lock(p);
          /*..update scheduling statistics .. */
          /* Possibly: current->thread_info.flags|=TIF_NEED_RESCHED */
          activate_task(rq,p,1);                 // Mark task as runnable
          p->state = TASK_RUNNING;
}
```

This code iterates over the wait list. For each task, it calls the `func` function. Most often, this function was set to `autoremove_wake_function` by `DEFINE_WAIT`, which in turn calls `try_to_wake_up`. Despite the name, this function not only tries but generally succeeds at waking the task up, by setting its state to `TASK_RUNNING` and placing it on a run queue. Again, names are misleading. The state `TASK_RUNNING` is really a READY state. At some later time, the task will be selected by the scheduler and will become the `current` running task. (The `.func` element of the wait queue entry is there as a "hook" to allow situation-specific code to be executed when the wakeup takes

place.)

After the task has been woken up, `autoremove_wake_function` then removes the task from the wait list, still under the protection of the wait queue's spin lock mutex. In the case of an exclusive wait, no further tasks are woken up, but otherwise the list iteration continues and additional tasks are awakened.

One or more of the awakened tasks might have a (dynamic) priority greater than the current task (the waker-upper). If this is true, `try_to_wakeup` will set the TIF_NEED_RESCHED flag for the current task, which will be checked as the task returns to user mode, causing schedule() to be called. Thus a newly awakened task may pre-empt the current task upon return to user mode.

A system call with blocking

As a further example, we shall follow a system call which may involve the caller being put to sleep, specifically reading from a pipe which currently has no data in it. In this example, the generic __wait_event macro is not used, but the kernel code effectively does the same thing. The Linux kernel is generally inconsistent, and depending on when a particular part of the kernel was coded or revised, different synchronization mechanisms may be in play.

As discussed in the previous unit, the parameters to the system call (file descriptor, buffer address, count) are passed via registers. At the entry to `system_call`, the registers are saved on the kernel mode stack, in the same order in which arguments are normally pushed on the stack in a regular C program. Therefore, when the particular `sys_xxx` system call handler is invoked, it finds the parameters on the stack just as if it had been invoked as an ordinary function. The `__user` macro in the argument declaration for the read buffer pointer is there to remind us that the address is a user-mode address and is not trusted. We pick up our system call trace at `fs/read_write.c:sys_read`:

```
// From /usr/src/linux/fs/read_write.c

asmlinkage ssize_t sys_read(unsigned int fd, char __user * buf, size_t count)
{
          struct file *file;
          ssize_t ret ;
          loff_t pos;

                    // Next 3 lines are interpolation of several routines
          if (fd >= current->files->max_fds) return -EBADF;
          if (!(file=current->files->fdtable[fd])) return -EBADF;
          file->f_count++
          pos=file->f_pos;
          ret = vfs_read(file, buf, count, &pos);
          file->f_pos=pos;
          file->f_count--;
```

```
          return ret;
}

ssize_t vfs_read(struct file *file, char __user *buf, size_t count, loff_t *pos)
{
        ssize_t ret;

        if (!(file->f_mode & FMODE_READ))               //open mode correct
                return -EBADF;
        if (!access_ok(VERIFY_WRITE, buf, count)))      //ptr to valid addr
                return -EFAULT;

          // rw_verify_area checks if the read request falls within
          // a mandatory record locking area of the file.  It also
          // does some range checks (e.g. negative position)
        ret = rw_verify_area(READ, file, pos, count);
          if (ret<0) return ret;
          count=ret;
          /*Most filesystems do not define op->read, but instead let the generic
            do_sync_read do the work by paging in the requested parts of file */
          if (file->f_op->read)
                    ret = file->f_op->read(file, buf, count, pos);  //VFS dispatch
          else
                    ret= do_sync_read(file, buf, count, pos);
          if (ret > 0) {
                    fsnotify_access(file->f_dentry);      //Hook for events
                    current->rchar += ret;                // Update stats
          }
          current->syscr++;                     // Update stats
        return ret;
}
```

As usual, the above code has been simplified somewhat from the actual Linux kernel sources. In particular, some complicated locking has been elided. The first thing done is to retrieve the `struct file` corresponding to the file descriptor. If the file descriptor is not valid, -EBADF is returned. The f_count field is incremented because an operation will be pending. Next, in the function `vfs_read`, other basic checks are performed. Was the file opened for reading (O_RDONLY or O_RDWR)? Is the read offset negative?

Finally, we are dispatched to the `read` method of the filesystem module which controls the file in question. This is performed via the `f_ops` structure of the `struct file`. Recall that the overall filesystem comprises one or more "mounted volumes," each of which may be of a different filesystem type. Each filesystem type has an associated module which provides methods for performing file operations such as read and write. (Un-named) pipes do not exist in the filesystem namespace, so there is a pseudo-filesystem module called `pipefs` which provides these methods when using pipe inodes. Our code walk-through winds up at `fs/pipe.c:pipe_read`:

```
ssize_t pipe_read(struct file *filp, char *buf, size_t count, loff_t *ppos)
{
        struct inode *inode = filp->f_dentry->d_inode;
```

```
        struct pipe_inode_info *info;
        ssize_t ret;
          int do_wakeup;

          do_wakeup=0;
          if (count==0) return 0;
          /* mutex_lock is a non-interruptible, blocking kernel mutex */
          mutex_lock(&inode->i_mutex);          /* Lock inode */
          info=inode->i_pipe;          /* Get private data */
          for(;;)
          {
            int bufs=info->nrbufs;
                    if (bufs > 0) {
                             /* ... This elided code copies the data from
                                      the pipe buffer(s) to user space and
                                      adjusts info->nrbufs accordingly.
                                      ret=number of bytes copied */
                             do_wakeup=1;                /* we have created room */
                             break;
                    }
                    else
                    {
                             /* Check for EOF condition (no writers ) */
                             if (info->writers==0)
                             {
                                      ret=0;
                                      break;
                             }
                             /* We have not read any data yet, so there is
                                no way that we need to wake up a blocked writer,
                                OK to sleep */
                             pipe_wait(inode);  //Releases and re-acquires mutex
                    }
                    if (signal_pending(current)) // Wait interrupted by sig
                    {
                              /* Only return ERESTARTSYS if we have read 0 bytes
                               * before waking up.  See text */
                             if (!ret) ret=-ERESTARTSYS;
                             break;
                    }
          }
          mutex_unlock(&inode->i_mutex);                  /* Release inode lock */
          if (do_wakeup) wake_up_interruptible_sync(&inode->i_pipe->wait);
          if (ret>0) file_access(filp);                   /* update atime*/
          return ret;
}

/* The arguments to pipe_write look strange because all writes within
          the kernel are transformed into a "scatter/gather" form known
          as an "iovec" with nr_segs segments.  */
ssize_t pipe_write(struct file *filp, struct iovec *_iov,
                    unsigned long nr_segs, loff_t pos)
{
```

```
  ssize_t chars;
  size_t total_len,ret;
  struct pipe_inode_info *pipe;
  struct inode *inode=filp->f_dentry->d_inode;
  int do_wakeup;
          total_len=iov_length(iov,nr_segs);    //Compute total write req len
          do_wakeup=0;
          ret=0;
          mutex_lock(&inode->mutex);
          pipe=inode->i_pipe;
          if (!pipe->readers) {
                    send_sig(SIGPIPE,current,0); // Unit 4
                    ret= -EPIPE;
                    goto out;
          }
          for(;;) {
                    bufs=pipe->nrbufs; //How many buffers already in pipe
                    if (bufs<PIPE_BUFFERS)      {         //still room
                              /* Elided code creates a new buffer, copies the
                                        user data into it, and puts buffer into
                                        the FIFO list of buffers */
                              do_wakeup=1;
                              ret += chars;
                              total_len -= chars;
                              if (!total_len) break;      //done copying
                    }
                    //No more room
                    //eliding code for NONBLOCK operation
                    /*If signal woke us up and we came back around the
                              loop, still full, do an interrupted system call */
                    if (signal_pending(current)) {
                      /* If we have written some bytes, return that value.
                         This is a case where write to a pipe CAN return
                         a "short write"!.  Do interrupted syscall only
                         if we hadn't written any bytes before being interrupted */
                              if (!ret) ret= -ERESTARTSYS;
                              break;
                    }
                    if (do_wakeup) {
                              /* We have written at least some data, make sure
                                 to wake up readers blocked on empty pipe, BEFORE
                                 sleeping on full pipe condition */
                              wake_up_interruptible_sync(&pipe_wait);
                              do_wakeup=0;
                    }
                    pipe->waiting_writers++;
                    pipe_wait(pipe);            //Wait for more room
                    pipe->waiting_writers--;
          }
 out:
          mutex_unlock(&inode->i_mutex);
          if (do_wakeup)  wake_up_interruptible_sync(&pipe->wait);
          if (ret>0) file_update_time(filp);    //Update inode mtime
```

```
        return ret;
}



void pipe_wait(struct inode *inode)
{
        DEFINE_WAIT(wait);          /* declare wait entry */
        prepare_to_wait(&inode->i_pipe->wait,&wait,TASK_INTERRUPTIBLE);
        mutex_unlock(&inode->i_mutex);
        schedule();
        finish_wait(&inode->i_pipe->wait,&wait);
        mutex_lock(&inode->i_mutex);
}
```

There are two sleep/wakeup conditions with which to contend: reader sleeping until more bytes are written, and writer sleeping until there is space in the pipe buffer. Let us consider the first case only, as the second case is analogous.

The first thing which `pipe_read` does is obtain a blocking mutex lock on the in-core inode. This will prevent other operations such as read, write, stat, etc. on the inode while the pipe_read is running. Because we are in a system call, which is a synchronous situation, it is acceptable to use this potentially blocking operation.

If there are buffered data waiting on the pipe, they are copied to the user's buffer (we are not concerned with the mechanics in this unit) and `do_wakeup` is set, which will remind us later to wake up potential waiting writers.

If there are no buffered data, and there are still possible writers to the pipe, then the calling process must be put to sleep. An INTERRUPTIBLE sleep is chosen because it could be a long time before a writer process wakes us up, and the user should have the option of breaking out of the sleep with a signal. The sleep will also be non-exclusive, because once data are placed into the pipe, it is meaningful for multiple readers to proceed (they might all want just a little data and there is enough to go around).

`pipe_wait` handles the business of going to sleep. A wait queue entry is defined as a local variable with `DEFINE_WAIT` and placed on the inode's wait list with `prepare_to_wait`. Now the inode mutex can be released (and indeed *must* be released, to avoid sleeping with the mutex held, which would prevent any writers from ever accessing the inode). `schedule` is called, and one or more other tasks run. Note that this code does not use the generic `__wait_event` code seen above but instead rolls its own. This is fairly typical of the non-uniform coding style of the Linux kernel.

Note that the inode mutex is held while the wait queue entry is made. This protects against the "lost wakeup" problem although the locking is fairly coarse -- only one reader or writer can ever be working on this pipe inode at any given time.

At some later time, another task writes to the pipe. This writer task will call, in `pipe_write()` `wakeup_interruptible_sync`, which eventually brings us into

the scheduler at `__wake_up_common`, seen above.

Therefore the reader process is eventually re-scheduled, and returns from pipe_wait(). Note the enclosing for(;;) loop which re-tests the condition (are there any data in the pipe?), because waking up does not necessarily mean the condition is now true.

Also note the check, after waking up, for pending signals. Posting a signal (which is not masked) to the waiting process here will cause the process to wake up, because the sleep is INTERRUPTIBLE. If some characters had already been read from the pipe before we had to sleep, then the read system call needs to return the number of characters read. However, if no characters were read, the system call will return ERESTARTSYS. See below under "Interrupted System Calls"

The reader (of these lecture notes, not the pipe) can also examine the pipe_read/pipe_write code above to see how a writer task, upon encountering a full pipe, goes to sleep and is awoken by a reader when the pipe is drained. Note how the pipe write atomicity (see Unit 4) is handled. Because pipe_write holds the mutex, no other writer's data can be interleaved. But, if there is not at least one full page-sized buffer available, or if the amount of data to write exceeds the available space in the pipe, the system call blocks after writing what it can. It is not an EXCLUSIVE wait so multiple writers can be blocked on the same pipe. When the pipe has room, they will all wake up but only one writer will win the mutex, and will get to write at least a 4K chunk of data atomically.

Sleep/Wakeup Summary

To summarize the sleeping and waking up process:
• Both the sleeper's code and the waker's code must agree on how the event is defined, and must both share a pre-declared wait queue.
SLEEPER:
• In the synchronous context of a system call or exception handler, allocates and initializes a wait queue entry
• The wait queue entry is added to the wait queue representing the event
• The task state is marked as TASK_INTERRUPTIBLE if the wait can be woken up by signal arrival, otherwise TASK_UNINTERRUPTIBLE.
• `schedule()` is called. The task is switched out, and because it is not in a runnable state, it will not get scheduled in again, until:
WAKER:
• The event arrives, either in a synchronous or an asynchronous context.
• The sleeping task is selected from the wait queue, and the wait queue entry is removed. For exclusive waits, only one task is selected, but for non-exclusive waits all waiting tasks are selected and removed from the wait queue.
• The task state is set to TASK_RUNNING and the task is placed on a run queue, making it eligible to be scheduled.

• If the woken task now has a "better" scheduling priority than the currently running task, the `TIF_NEED_RESCHED` flag is set on the current task, which subjects it to pre-emption by the woken task when execution returns to user level from the system call, fault or interrupt.
• The woken task is (at some later time) selected by the scheduler to run, and is switched in.

Interrupted System Calls / System Call Restart

In Linux kernel system call code, when a sleeping process is awoken not because the event has arrived, but because a signal is posted to that process (and that signal is not masked), that is known as an Interrupted System Call. This is a poorly documented aspect of UNIX systems programming. From the standpoint of the system call code, it can not continue because the event has not arrived and the signal needs to be delivered, therefore we need to return out of the system call and head back towards userland, at which point the signal can be delivered. The system call code returns one of the following error codes:
• EINTR: This system call can not be restarted. The error EINTR will be returned to the user.
• ERESTARTSYS: The system call will be restarted if the disposition of the signal in question is DFL or IGN. (Note: a signal which is being IGNored but which isn't masked will still cause a sleep to be woken up. Also note that while the DeFauLt action for many signals is process termination, whereupon system call restart is a moot point, there are a few where the default action is not termination). If there is a handler defined, the system call will only by restarted if the SA_RESTART flag is set for that signal number (e.g. through the `sigaction` system call). System call restart will happen after the handler returns. Most system calls use ERESTARTSYS.
• ERESTARTNOINTR: The system call will always be restarted, even if the handler does not have SA_RESTART set.
• ERESTARTNOHAND: The system call will be restarted if the signal disposition is DFL or IGN. If there is a handler, the system call will not be restarted, regardless of SA_RESTART, and upon completion of the handler EINTR will be returned from the system call.
• ERESTART_RESTARTBLOCK: This is a special case used for a few time-related system calls which require specific code to be executed prior to system call restart to adjust those time-related parameters. It forces a special `restart_syscall` system call to be executed upon resumption in userland to make this adjustments.

System call restart is effected by adjusting the user-mode registers prior to returning to user mode. The eip register (program counter) is decremented so that upon resumption of userland code, the original `INT $0x80` opcode is executed again. The eax is reset to the original system call value. This is accomplished by modifying the saved user-mode

registers which are on the kernel stack (see previous unit). Upon resumption of the user-mode program (and return from the signal handler if applicable) the system call will be re-called with the exact same parameters.

Process Termination

A process terminates either when it has received a fatal signal or it explicitly calls the `exit` system call. Let us look at the exit system call, ignoring some of the complexities introduced by threads and thread groups. `sys_exit` is a simple wrapper for `do_exit`. The latter can also be called during signal delivery (see unit 11).

```
asmlinkage long sys_exit(int error_code)
{
        do_exit((error_code&0xff)<<8);
}


void do_exit(long code)
{
        struct task_struct *tsk = current;

                    /* These are "can't happens" */
        if (unlikely(in_interrupt()))
                panic("Aiee, killing interrupt handler!");
        if (unlikely(tsk->pid==0))
                panic("Attempted to kill the idle task!");
        if (unlikely(tsk->pid == 1))
                panic("Attempted to kill init!");
          tracehook_report_exit(&code);         /* Hook for strace */
        tsk->flags |= PF_EXITING;
        exit_mm(tsk);                   /* Release virtual address space */
          exit_signals(tsk);  /* Cleanup/reassign pending signals */
        exit_files(tsk);     /* Close open files */
        exit_fs(tsk);                   /* Leave cwd, etc. */
        exit_thread();                  /* Misc. cleanup */
          tsk->exit_code = code;        /* Exit cause code for wait() syscall */
          exit_notify(tsk);  /* Send signal to and/or wake-up parent */
        tsk->state= TASK_DEAD;
        schedule();
        BUG();              /* If control reaches here we are in trouble! */
}
```

Each of the helper routines can have further consequences. E.g. when `__exit_files` closes each file descriptor, if that drops the references to the `struct file` to 0, then that structure is de-allocated. If in turn there are no other `struct file` instances in the system pointing to the in-core inode, then the inode is closed. If the inode had been unlinked while open, its resources are freed, etc. After sending SIGCHLD to the parent process, `exit_notify` places the exiting process in the `EXIT_ZOMBIE` state. Although most of the resources have been released, the `struct task` remains to hold

the exit code and other statistics of the process's life, such as cpu time accumulated, for collection by a parent with one of the many variants of the `wait` system call. After this is done, the `struct task` is finally released and the process id is available for recycling. The zombie task is never scheduled, because nothing puts it onto a scheduler run queue. If somehow the last line is reached, it is evidence of a kernel bug!