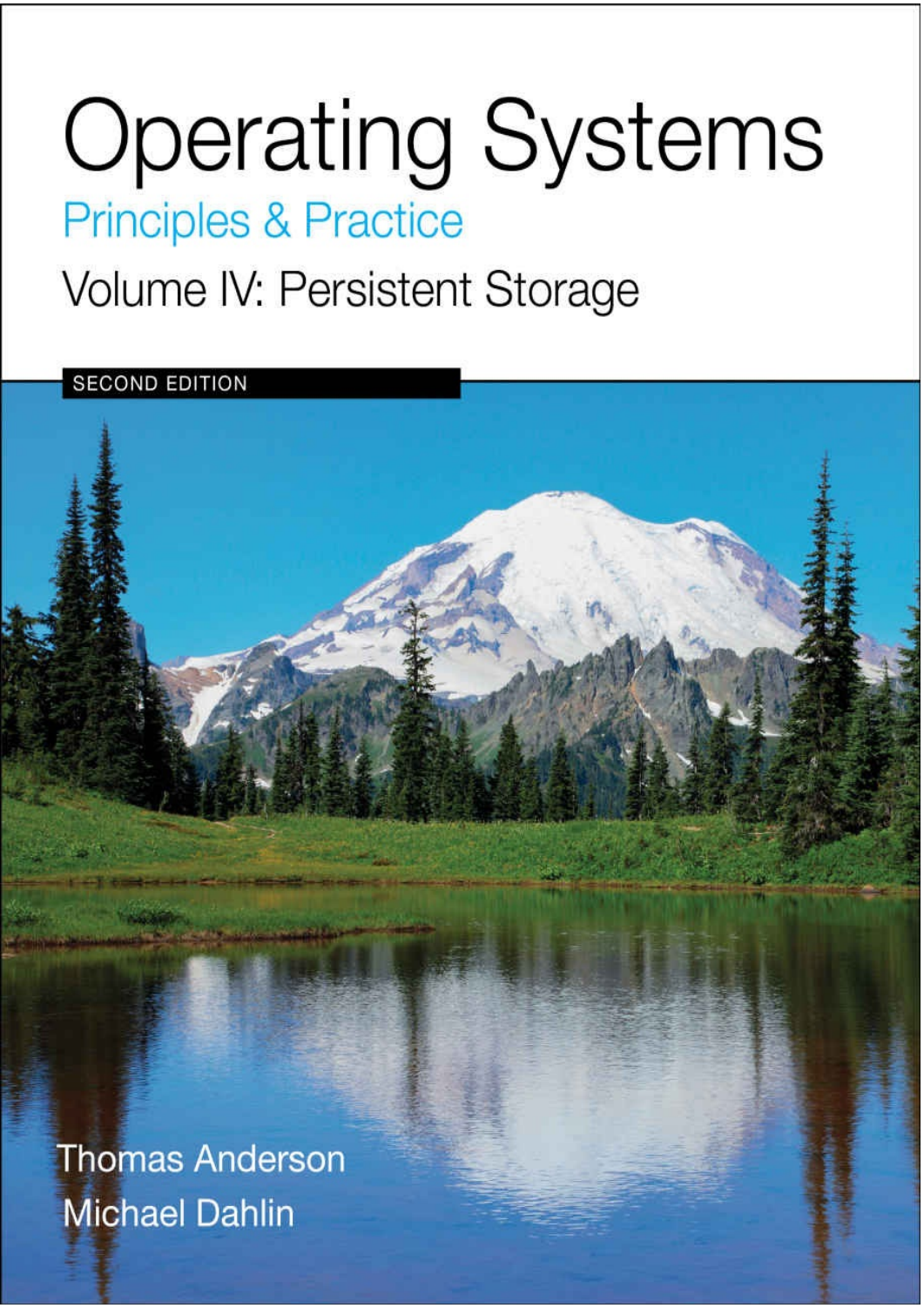


Operating Systems

Principles & Practice

Volume IV: Persistent Storage

Second Edition

Thomas Anderson

University of Washington

Mike Dahlin

University of Texas and Google

Recursive Books

recursivebooks.com

Operating Systems: Principles and Practice (Second Edition) Volume IV: Persistent Storage by Thomas Anderson and Michael Dahlin

ISBN 978-0-9856735-6-7
Publisher: Recursive Books, Ltd., http://recursivebooks.com/
Cover: Reflection Lake, Mt. Rainier
Cover design: Cameron Neat
Illustrations: Cameron Neat
Copy editors: Sandy Kaplan, Whitney Schmidt
Ebook design: Robin Briggs
Web design: Adam Anderson

SUGGESTIONS, COMMENTS, and ERRORS. We welcome suggestions, comments and error reports, by email to suggestions@recursivebooks.com

To Robin, Sandra, Katya, and Adam
Tom Anderson

To Marla, Kelly, and Keith
Mike Dahlin

Contents

Preface

Preface to the eBook Edition

Operating Systems: Principles and Practice is a textbook for a first course in undergraduate operating systems. In use at over 50 colleges and universities worldwide, this textbook provides:

- A path for students to understand high level concepts all the way down to working code.
- Extensive worked examples integrated throughout the text provide students concrete guidance for completing homework assignments.
- A focus on up-to-date industry technologies and practice

The eBook edition is split into four volumes that together contain exactly the same material as the (2nd) print edition of Operating Systems: Principles and Practice, reformatted for various screen sizes. Each volume is self-contained and can be used as a standalone text, e.g., at schools that teach operating systems topics across multiple courses.

- **Volume 1: Kernels and Processes.** This volume contains Chapters 1-3 of the print edition. We describe the essential steps needed to isolate programs to prevent buggy applications and computer viruses from crashing or taking control of your system.
- **Volume 2: Concurrency.** This volume contains Chapters 4-7 of the print edition. We provide a concrete methodology for writing correct concurrent programs that is in widespread use in industry, and we explain the mechanisms for context switching and synchronization from fundamental concepts down to assembly code.
- **Volume 3: Memory Management.** This volume contains Chapters 8-10 of the print edition. We explain both the theory and mechanisms behind 64-bit address space translation, demand paging, and virtual machines.
- **Volume 4: Persistent Storage.** This volume contains Chapters 11-14 of the print edition. We explain the technologies underlying modern extent-based, journaling, and versioning file systems.

A more detailed description of each chapter is given in the preface to the print edition.

Preface to the Print Edition

Why We Wrote This Book

Many of our students tell us that operating systems was the best course they took as an undergraduate and also the most important for their careers. We are not alone — many of our colleagues report receiving similar feedback from their students.

Part of the excitement is that the core ideas in a modern operating system — protection, concurrency, virtualization, resource allocation, and reliable storage — have become widely applied throughout computer science, not just operating system kernels. Whether you get a job at Facebook, Google, Microsoft, or any other leading-edge technology company, it is impossible to build resilient, secure, and flexible computer systems without the ability to apply operating systems concepts in a variety of settings. In a modern world, nearly everything a user does is distributed, nearly every computer is multi-core, security threats abound, and many applications such as web browsers have become mini-operating systems in their own right.

It should be no surprise that for many computer science students, an undergraduate operating systems class has become a *de facto* requirement: a ticket to an internship and eventually to a full-time position.

Unfortunately, many operating systems textbooks are still stuck in the past, failing to keep pace with rapid technological change. Several widely-used books were initially written in the mid-1980's, and they often act as if technology stopped at that point. Even when new topics are added, they are treated as an afterthought, without pruning material that has become less important. The result are textbooks that are very long, very expensive, and yet fail to provide students more than a superficial understanding of the material.

Our view is that operating systems have changed dramatically over the past twenty years, and that justifies a fresh look at both *how* the material is taught and *what* is taught. The pace of innovation in operating systems has, if anything, increased over the past few years, with the introduction of the iOS and Android operating systems for smartphones, the shift to multicore computers, and the advent of cloud computing.

To prepare students for this new world, we believe students need three things to succeed at understanding operating systems at a deep level:

- **Concepts and code.** We believe it is important to teach students both *principles* and *practice*, concepts and implementation, rather than either alone. This textbook takes concepts all the way down to the level of working code, e.g., how a context switch works in assembly code. In our experience, this is the only way students will really understand and master the material. All of the code in this book is available from the author's web site, ospp.washington.edu.
- **Extensive worked examples.** In our view, students need to be able to apply concepts in practice. To that end, we have integrated a large number of example exercises, along with solutions, throughout the text. We uses these exercises extensively in our own lectures, and we have found them essential to challenging students to go beyond a superficial understanding.
- **Industry practice.** To show students how to apply operating systems concepts in a variety of settings, we use detailed, concrete examples from Facebook, Google, Microsoft, Apple, and other leading-edge technology companies throughout the textbook. Because operating systems concepts are important in a wide range of computer systems, we take these examples not only from traditional operating systems like Linux, Windows, and OS X but also from other systems that need to solve problems of protection, concurrency, virtualization, resource allocation, and reliable storage like databases, web browsers, web servers, mobile applications, and search engines.

Taking a fresh perspective on what students need to know to apply operating systems concepts in practice has led us to innovate in every major topic covered in an undergraduate-level course:

- **Kernels and Processes.** The safe execution of untrusted code has become central to many types of computer systems, from web browsers to virtual machines to operating systems. Yet existing textbooks treat protection as a side effect of UNIX processes, as if they are synonyms. Instead, we start from first principles: what are the minimum requirements for process isolation, how can systems implement process isolation efficiently, and what do students need to know to implement functions correctly when the caller is potentially malicious?
- **Concurrency.** With the advent of multi-core architectures, most students today will spend much of their careers writing concurrent code. Existing textbooks provide a blizzard of concurrency alternatives, most of which were abandoned decades ago as impractical. Instead, we focus on providing students a *single* methodology based on Mesa monitors that will enable students to write correct concurrent programs — a methodology that is by far the dominant approach used in industry.
- **Memory Management.** Even as demand-paging has become less important, virtualization has become even more important to modern computer systems. We provide a deep treatment of address translation hardware, sparse address spaces, TLBs, and on-chip caches. We then use those concepts as a springboard for describing virtual machines and related concepts such as checkpointing and copy-on-write.
- **Persistent Storage.** Reliable storage in the presence of failures is central to the design of most computer systems. Existing textbooks survey the history of file systems, spending most of their time ad hoc approaches to failure recovery and de-fragmentation. Yet no modern file systems still use those ad hoc approaches. Instead, our focus is on how file systems use extents, journaling, copy-on-write, and RAID to achieve both high performance and high reliability.

Intended Audience

Operating Systems: Principles and Practice is a textbook for a first course in undergraduate operating systems. We believe operating systems should be taken as early as possible in an undergraduate's course of study; many students use the course as a springboard to an internship and a career. To that end, we have designed the textbook to assume minimal pre-requisites: specifically, students should have taken a data structures course and one on computer organization. The code examples are written in a combination of x86 assembly, C, and C++. In particular, we have designed the book to interface well with the Bryant and O'Halloran textbook. We review and cover in much more depth the material from the second half of that book.

We should note what this textbook is *not*: it is not intended to teach the API or internals of any specific operating system, such as Linux, Android, Windows 8, OS X, or iOS. We use many concrete examples from these systems, but our focus is on the shared problems these systems face and the technologies these systems use to solve those problems.

A Guide to Instructors

One of our goals is enable instructors to choose an appropriate level of depth for each course topic. Each chapter begins at a conceptual level, with implementation details and the more advanced material towards the end. The more advanced material can be omitted without compromising the ability of students to follow later material. No single-quarter or single-semester course is likely to be able to cover every topic we have included, but we think it is a good thing for students to come away from an operating systems course with an appreciation that there is *always* more to learn.

For each topic, we attempt to convey it at three levels:

- **How to reason about systems.** We describe core systems concepts, such as protection, concurrency, resource scheduling, virtualization, and storage, and we provide practice applying these concepts in various situations. In our view, this provides the biggest long-term payoff to students, as they are likely to need to apply these concepts in their work throughout their career, almost regardless of what project they end up working on.
- **Power tools.** We introduce students to a number of abstractions that they can apply in their work in industry immediately after graduation, and that we expect will continue to be useful for decades such as sandboxing, protected procedure calls, threads, locks, condition variables, caching, checkpointing, and transactions.
- **Details of specific operating systems.** We include numerous examples of how different operating systems work in practice. However, this material changes rapidly, and there is an order of magnitude more material than can be covered in a single semester-length course. The purpose of these examples is to illustrate how to use the operating systems principles and power tools to solve concrete problems. We do not attempt to provide a comprehensive description of Linux, OS X, or any other particular operating system.

The book is divided into five parts: an introduction (Chapter 1), kernels and processes (Chapters 2-3), concurrency, synchronization, and scheduling (Chapters 4-7), memory management (Chapters 8-10), and persistent storage (Chapters 11-14).

- **Introduction.** The goal of Chapter 1 is to introduce the recurring themes found in the later chapters. We define some common terms, and we provide a bit of the history of the development of operating systems.
- **The Kernel Abstraction.** Chapter 2 covers kernel-based process protection — the concept and implementation of executing a user program with restricted privileges. Given the increasing importance of computer security issues, we believe protected execution and safe transfer across privilege levels are worth treating in depth. We have broken the description into sections, to allow instructors to choose either a quick introduction to the concepts (up through Section 2.3), or a full treatment of the kernel implementation details down to the level of interrupt handlers. Some instructors start

with concurrency, and cover kernels and kernel protection afterwards. While our textbook can be used that way, we have found that students benefit from a basic understanding of the role of operating systems in executing user programs, before introducing concurrency.

- **The Programming Interface.** Chapter 3 is intended as an impedance match for students of differing backgrounds. Depending on student background, it can be skipped or covered in depth. The chapter covers the operating system from a programmer's perspective: process creation and management, device-independent input/output, interprocess communication, and network sockets. Our goal is that students should understand at a detailed level what happens when a user clicks a link in a web browser, as the request is transferred through operating system kernels and user space processes at the client, server, and back again. This chapter also covers the organization of the operating system itself: how device drivers and the hardware abstraction layer work in a modern operating system; the difference between a monolithic and a microkernel operating system; and how policy and mechanism are separated in modern operating systems.
- **Concurrency and Threads.** Chapter 4 motivates and explains the concept of threads. Because of the increasing importance of concurrent programming, and its integration with modern programming languages like Java, many students have been introduced to multi-threaded programming in an earlier class. This is a bit dangerous, as students at this stage are prone to writing programs with race conditions, problems that may or may not be discovered with testing. Thus, the goal of this chapter is to provide a solid conceptual framework for understanding the semantics of concurrency, as well as how concurrent threads are implemented in both the operating system kernel and in user-level libraries. Instructors needing to go more quickly can omit these implementation details.
- **Synchronization.** Chapter 5 discusses the synchronization of multi-threaded programs, a central part of all operating systems and increasingly important in many other contexts. Our approach is to describe one effective method for structuring concurrent programs (based on Mesa monitors), rather than to attempt to cover several different approaches. In our view, it is more important for students to master one methodology. Monitors are a particularly robust and simple one, capable of implementing most concurrent programs efficiently. The implementation of synchronization primitives should be included if there is time, so students see that there is no magic.
- **Multi-Object Synchronization.** Chapter 6 discusses advanced topics in concurrency — specifically, the twin challenges of multiprocessor lock contention and deadlock. This material is increasingly important for students working on multicore systems, but some courses may not have time to cover it in detail.
- **Scheduling.** This chapter covers the concepts of resource allocation in the specific context of processor scheduling. With the advent of data center computing and multicore architectures, the principles and practice of resource allocation have renewed importance. After a quick tour through the tradeoffs between response time and throughput for uniprocessor scheduling, the chapter covers a set of more advanced topics in affinity and multiprocessor scheduling, power-aware and deadline scheduling, as well as basic queueing theory and overload management. We conclude these topics by walking students through a case study of server-side load management.
- **Address Translation.** Chapter 8 explains mechanisms for hardware and software address translation. The first part of the chapter covers how hardware and operating systems cooperate to provide flexible, sparse address spaces through multi-level segmentation and paging. We then describe how to make memory management efficient with translation lookaside buffers (TLBs) and virtually addressed caches. We consider how to keep TLBs consistent when the operating system makes changes to its page tables. We conclude with a discussion of modern software-based protection mechanisms such as those found in the Microsoft Common Language Runtime and Google's Native Client.
- **Caching and Virtual Memory.** Caches are central to many different types of computer systems. Most students will have seen the concept of a cache in an earlier class on machine structures. Thus, our goal is to cover the theory and implementation of caches: when they work and when they do not, as well as how they are implemented in hardware and software. We then show how these ideas are applied in the context of memory-mapped files and demand-paged virtual memory.
- **Advanced Memory Management.** Address translation is a powerful tool in system design, and we show how it can be used for zero copy I/O, virtual machines, process checkpointing, and recoverable virtual memory. As this is more advanced material, it can be skipped by those classes pressed for time.
- **File Systems: Introduction and Overview.** Chapter 11 frames the file system portion of the book, starting top down with the challenges of providing a useful file abstraction to users. We then discuss the UNIX file system interface, the major internal elements inside a file system, and how disk device drivers are structured.
- **Storage Devices.** Chapter 12 surveys block storage hardware, specifically magnetic disks and flash memory. The last two decades have seen rapid change in storage technology affecting both application programmers and operating systems designers; this chapter provides a snapshot for students, as a building block for the next two chapters. If students have previously seen this material, this chapter can be skipped.
- **Files and Directories.** Chapter 13 discusses file system layout on disk. Rather than survey all possible file layouts — something that changes rapidly over time — we use file systems as a concrete example of mapping complex data structures onto block storage devices.
- **Reliable Storage.** Chapter 14 explains the concept and implementation of reliable storage, using file systems as a concrete example. Starting with the ad hoc techniques used in early file systems, the chapter explains checkpointing and write ahead logging as alternate implementation strategies for building reliable storage, and it discusses how redundancy such as checksums and replication are used to improve reliability and availability.

We welcome and encourage suggestions for how to improve the presentation of the material; please send any comments to the publisher's website, suggestions@recursivebooks.com.

Acknowledgements

We have been incredibly fortunate to have the help of a large number of people in the conception, writing, editing, and production of this book.

We started on the journey of writing this book over dinner at the USENIX NSDI conference in 2010. At the time, we thought perhaps it would take us the summer to complete the first version and perhaps a year before we could declare ourselves done. We were very wrong! It is no exaggeration to say that it would have taken us a lot longer without the help we have received from the people we mention below.

Perhaps most important have been our early adopters, who have given us enormously useful feedback as we have put together this edition:

Carnegie-Mellon	David Eckhardt and Garth Gibson
Clarkson	Jeanna Matthews
Cornell	Gun Sirer
ETH Zurich	Mothy Roscoe
New York University	Laskshmi Subramanian
Princeton University	Kai Li
Saarland University	Peter Druschel
Stanford University	John Ousterhout
University of California Riverside	Harsha Madhyastha
University of California Santa Barbara	Ben Zhao
University of Maryland	Neil Spring
University of Michigan	Pete Chen
University of Southern California	Ramesh Govindan
University of Texas-Austin	Lorenzo Alvisi
Universtiy of Toronto	Ding Yuan
University of Washington	Gary Kimura and Ed Lazowska

In developing our approach to teaching operating systems, both before we started writing and afterwards as we tried to put our thoughts to paper, we made extensive use of lecture notes and slides developed by other faculty. Of particular help were the materials created by Pete Chen, Peter Druschel, Steve Gribble, Eddie Kohler, John Ousterhout, Mothy Roscoe, and Geoff Voelker. We thank them all.

Our illustrator for the second edition, Cameron Neat, has been a joy to work with. We would also like to thank Simon Peter for running the multiprocessor experiments introducing Chapter 6.

We are also grateful to Lorenzo Alvisi, Adam Anderson, Pete Chen, Steve Gribble, Sam Hopkins, Ed Lazowska, Harsha Madhyastha, John Ousterhout, Mark Rich, Mothy Roscoe, Will Scott, Gun Sirer, Ion Stoica, Lakshmi Subramanian, and John Zahorjan for their helpful comments and suggestions as to how to improve the book.

We thank Josh Berlin, Marla Dahlin, Rasit Eskicioglu, Sandy Kaplan, John Ousterhout, Whitney Schmidt, and Mike Walfish for helping us identify and correct grammatical or technical bugs in the text.

We thank Jeff Dean, Garth Gibson, Mark Oskin, Simon Peter, Dave Probert, Amin Vahdat, and Mark Zbikowski for their help in explaining the internal workings of some of the commercial systems mentioned in this book.

We would like to thank Dave Wetherall, Dan Weld, Mike Walfish, Dave Patterson, Olav Kvern, Dan Halperin, Armando Fox, Robin Briggs, Katya Anderson, Sandra Anderson, Lorenzo Alvisi, and William Adams for their help and advice on textbook economics and production.

The Helen Riaboff Whiteley Center as well as Don and Jeanne Dahlin were kind enough to lend us a place to escape when we needed to get chapters written.

Finally, we thank our families, our colleagues, and our students for supporting us in this larger-than-expected effort.

IV
Persistent Storage

11. File Systems: Introduction and Overview

Memory is the treasury and guardian of all things. —*Marcus Tullius Cicero*

Computers must be able to reliably store data. Individuals store family photos, music files, and email folders; programmers store design documents and source files; office workers store spreadsheets, text documents, and presentation slides; and businesses store inventory, orders, and billing records. In fact, for a computer to work at all, it needs to be able to store programs to run and the operating system, itself.

For all of these cases, users demand a lot from their storage systems:

- **Reliability.** A user's data should be safely stored even if a machine's power is turned off or its operating system crashes. In fact, much of this data is so important that users expect and need the data to survive even if the devices used to store it are damaged. For example, many modern storage systems continue to work even if one of the magnetic disks storing the data malfunctions or even if a data center housing some of the system's servers burns down!
- **Large capacity and low cost.** Users and companies store enormous amount of data, so they want to be able to buy high capacity storage for a low cost. For example, it takes about 350 MB to store an hour of CD-quality losslessly encoded music, 4 GB to store an hour-long high-definition home video, and about 1 GB to store 300 digital photos. As a result of these needs, many individuals own 1 TB or more of storage for their personal files. This is an enormous amount: if you printed 1 TB of data as text on paper, you would produce a stack about 20 miles high. In contrast, for less than $100 you can buy 1 TB of storage that fits in a shoebox.
- **High performance.** For programs to use data, they must be able to access it, and for programs to use large amounts of data, this access must be fast. For example, users want program start-up to be nearly instantaneous, a business may need to process hundreds or thousands of orders per second, or a server may need to stream a large number of video files to different users.
- **Named data.** Because users store a large amount of data, because some data must last longer than the process that creates it, and because data must be shared across programs, storage systems must provide ways to easily identify data of interest. For example, if you can name a file (e.g., /home/alice/assignments/hw1.txt) you can find the data you want out of the millions of blocks on your disk, you can still find it after you shut down your text editor, and you can use your email program to send the data produced by the text editor to another user.
- **Controlled sharing.** Users need to be able to share stored data, but this sharing needs to be controlled. As one example, you may want to create a design document that everyone in your group can read and write, that people in your department can read but not write, and that people outside of your department cannot access at all. As another example, it is useful for a system to be able to allow anyone to execute a

program while only allowing the system administrator to change the program.

Nonvolatile storage and file systems. The contents of a system's main DRAM memory can be lost if there is an operating system crash or power failure. In contrast, *non-volatile storage* is durable and retains its state across crashes and power outages; non-volatile storage is also called or *persistent storage* or *stable storage*. Nonvolatile storage can also have much higher capacity and lower cost than the volatile DRAM that forms the bulk of most system's "main memory."

However, non-volatile storage technologies have their own limitations. For example, current non-volatile storage technologies such as magnetic disks and high-density flash storage do not allow random access to individual words of storage; instead, access must be done in more coarse-grained units — 512, 2048, or more bytes at a time.

Furthermore, these accesses can be much slower than access to DRAM; for example, reading a sector from a magnetic disk may require activating a motor to move a disk arm to a desired track on disk and then waiting for the spinning disk to bring the desired data under the disk head. Because disk accesses involve motors and physical motion, the time to access a random sector on a disk can be around 10 milliseconds. In contrast, DRAM latencies are typically under 100 nanoseconds. This large difference — about five orders of magnitude in the case of spinning disks — drives the operating system to organize and use persistent storage devices differently than main memory.

File systems are a common operating system abstraction to allow applications to access non-volatile storage. File systems use a number of techniques to cope with the physical limitations of non-volatile storage devices and to provide better abstractions to users. For example, Figure 11.1 summarizes how physical characteristics motivate several key aspects of file system design.

Goal	Physical Characteristic	Design Implication
High performance	Large cost to initiate IO access	Organize data placement with *files*, *directories*, *free space bitmap*, and *placement heuristics* so that storage is accessed in large sequential units
		Caching to avoid accessing persistent storage
Named data	Storage has large capacity, survives crashes, and is shared across programs	Support *files* and *directories* with meaningful names
Controlled sharing	Device stores many users' data	Include access-control *metadata* with files
Reliable storage	Crash can occur during update	Use *transactions* to make a set of updates atomic
	Storage devices can fail	Use *redundancy* to detect and correct failures
	Flash memory cells can wear out	Move data to different storage locations to even the wear

Figure 11.1: Characteristics of persistent storage devices affect the design of an operating system's storage abstractions.

- **Performance.** File systems amortize the cost of initiating expensive operations — such as moving a disk arm or erasing a block of solid state memory — by grouping where its placement of data so that such operations access large, sequential ranges of storage.
- **Naming.** File systems group related data together into directories and files and provide human-readable names for them (e.g., /home/alice/Pictures/summer-vacation/hiking.jpg.) These names for data remain meaningful even after the program that creates the data exits, they help users organize large amounts of storage, and they make it easy for users to use different programs to create, read, and edit, their data.
- **Controlled sharing.** File systems include metadata about who owns which files and which other users are allowed to read, write, or execute data and program files.
- **Reliability.** File systems use transactions to atomically update multiple blocks of persistent storage, similar to how the operating system uses critical sections to atomically update different data structures in memory.

  To further improve reliability, file systems store checksums with data to detect corrupted blocks, and they replicate data across multiple storage devices to recover from hardware failures.

Impact on application writers. Understanding the reliability and performance properties of storage hardware and file systems is important even if you are not designing a file system from scratch. Because of the fundamental limitations of existing storage devices, the higher-level illusions of reliability and performance provided by the file system are imperfect. An application programmer needs to understand these limitations to avoid having inconsistent data stored on disk or having a program run orders of magnitude slower than expected.

For example, suppose you edit a large document with many embedded images and that your word processor periodically auto-saves the document so that you would not lose too many edits if the machine crashes. If the application uses the file system in a straightforward way, several of unexpected things may happen.

- **Poor performance.** First, although file systems allow existing bytes in a file to be overwritten with new values, they do not allow new bytes to be inserted into the middle of existing bytes. So, even a small update to the file may require rewriting the entire file either from beginning to end or at least from the point of the first insertion to the end. For a multi-megabyte file, each auto-save may end up taking as much as a second.
- **Corrupt file.** Second, if the application simply overwrites the existing file with updated data, an untimely crash can leave the file in an inconsistent state, containing a mishmash of the old and new versions. For example, if a section is cut from one location and pasted in another, after a crash the saved document may end up with copies of the section in both locations, one location, or neither location; or it may end up with a region that is a mix of the old and new text.
- **Lost file.** Third, if instead of overwriting the document file, the application writes updates to a new file, then deletes the original file, and finally moves the new file to the original file's location, an untimely crash can leave the system with no copies of the document at all.

Programs use a range of techniques to deal with these types of issues. For example, some structure their code to take advantage of the detailed semantics of specific operating systems. Some operating systems guarantee that when a file is renamed and a file with the target name already exists, the target name will always refer to either the old or new file, even after a crash in the middle of the rename operation. In such a case, an implementation can create a new file with the new version of the data and use the rename command to atomically replace the old version with the new one.

Other programs essentially build a miniature file system over the top of the underlying one, structuring their data so that the underlying file system can better meet their performance and reliability requirements.

For example, a word processor might use a sophisticated document format, allowing it to, for example, add and remove embedded images and to always update a document by appending updates to the end of the file.

As another example, a data analysis program might improve its performance by organizing its accesses to input files in a way that ensures that each input file is read only once and that it is read sequentially from its start to its end.

Or, a browser with a 1 GB on-disk cache might create 100 files, each containing 10 MB of data, and group a given web site's objects in a sequential region of a randomly selected file. To do this, the browser would need to keep metadata that maps each cached web site to a region of a file, it would need to keep track of what regions of each file are used and which are free, it would need to decide where to place a new web site's objects, and it would need to have a strategy for growing or moving a web site's objects as additional objects are fetched.

Roadmap. To get good performance and acceptable reliability, both application writers and operating systems designers must understand how storage devices and file systems work. This chapter and the next three discuss the key issues:

- **API and abstractions.** The rest of this chapter introduces file systems by describing a typical API and set of abstractions, and it provides an overview of the software layers that provide these abstractions.
- **Storage devices.** The characteristics of persistent storage devices strongly influence the design of storage system abstractions and higher level applications. Chapter 12 therefore explores the physical characteristics of common storage devices.
- **Implementing files and directories.** Chapter 13 describes how file systems keep track of data by describing several widely used approaches to implementing files and directories.
- **Reliable storage.** Although we would like storage to be perfectly reliable, physical devices fall short of that ideal. Chapter 14 describes how storage systems use transactional updates and redundancy to improve reliability.

11.1 The File System Abstraction

Today, almost anyone who uses a computer is familiar with the high-level file system abstraction. File systems provide a way for users to organize their data and to store it for long periods of time. For example, Bob's computer might store a collection of applications such as /Applications/Calculator and /Program Files/Text Edit and a collection of data files such as /home/Bob/correspondence/letter-to-mom.txt, and /home/Bob/Classes/OS/hw1.txt.

More precisely, a *file system* is an operating system abstraction that provides persistent, named data. *Persistent data* is stored until it is explicitly deleted, even if the computer storing it crashes or loses power. *Named data* can be accessed via a human-readable identifier that the file system associates with the file. Having a name allows a file to be accessed even after the program that created it has exited, and allows it to be shared by multiple applications.

There are two key parts to the file system abstraction: files, which define sets of data, and directories, which define names for files.

File. A *file* is a named collection of data in a file system. For example, the programs /Applications/Calculator or /Program Files/Text Edit are each files, as are the data /home/Bob/correspondence/letter-to-mom.txt or /home/Bob/Classes/OS/hw1.txt.

Files provide a higher-level abstraction than the underlying storage device: they let a single, meaningful name refer to an (almost) arbitrarily-sized amount of data. For example /home/Bob/Classes/OS/hw1.txt might be stored on disk in blocks 0x0A713F28, 0xB3CA349A, and 0x33A229B8, but it is much more convenient to refer to the data by its name than by this list of disk addresses.

A file's information has two parts, metadata and data. A *file's metadata* is information about the file that is understood and managed by the operating system. For example, a file's metadata typically includes the file's *size*, its *modification time*, its *owner*, and its *security information* such as whether it may be read, written, or executed by the owner or by other users.

A *file's data* can be whatever information a user or application puts in it. From the point of

view of the file system, a file's data is just an array of untyped bytes. Applications can use these bytes to store whatever information they want in whatever format they choose. Some data have a simple structure. For example, an ASCII text file contains a sequence of bytes that are interpreted as letters in the English alphabet. Conversely, data structures stored by applications can be arbitrarily complex. For example, a .doc files can contain text, formatting information, and embedded objects and images, an ELF (Executable and Linkable File) files can contain compiled objects and executable code, or a database file can contain the information and indices managed by a relational database.

Executing "untyped" files

Usually, an operating system treats a file's data as an array of untyped bytes, leaving it up to applications to interpret a file's contents. Occasionally, however, the operating system needs to be able to parse a file's data.

For example, Linux supports a number of different executable file types such as the ELF and a.out binary files and tcsh, csh, and perl scripts. You can run any of these files from the command line or using the exec() system call. E.g.,
> a.out
Hello world from hello.c compiled by gcc!
> hello.pl
Hello world from hello.pl, a perl script!
> echo ''Hello world from /bin/echo, a system binary!''
Hello world from /bin/echo, a system binary!

To execute a file, the operating system must determine whether it is a binary file or a script. If it is the former, the operating system must parse the file to determine where in the target process's memory to load code and data from the file and which instruction to start with. If it is the latter, the operating system must determine which interpreter program it should launch to execute the script.

Linux does this by having executable files begin with a *magic number* that identifies the file's format. For example, ELF binary executables begin with the four bytes 0x7f, 0x45, 0x4c, and 0x46 (the ASCII characters DEL, E, L, and F); once an executable is known to be an ELF file, the ELF standard defines how the operating system should parse the rest of the file to extract and load the program's code and data. Similarly, script files begin with #! followed by the name of the interpreter that should be used to run the script (e.g., a script might begin with #! /bin/sh to be executed using the Bourne shell or #! /usr/bin/perl to be executed using the perl interpreter.

Alternative approaches include determining a file's type by its name *extension* — the characters after the last dot (.) in the file's name (e.g., .exe, .pl, or .sh) — or including information about a file's type in its metadata.

Multiple data streams

For traditional files, the file's data is a single logical sequence of bytes, and each byte can be identified by its offset from the start of the sequence (e.g., byte 0, byte 999, or byte 12481921 of a file.)

Some file systems support multiple sequences of bytes per file. For example, Apple's MacOS Extended file system supports multiple *forks* per file — a data fork for the file's basic data, a resource fork for storing additional attributes for the file, and multiple named forks for application-defined data. Similarly, Microsoft's NTFS supports *alternate data streams* that are similar to MacOS's named forks.

In these systems, when you open a file to read or write its data, you specify not only the file but also the fork or stream you want.

Directory. Whereas a file contains system-defined metadata and arbitrary data, directories provide names for files. In particular, a *file directory* is a list of human-readable names and a mapping from each name to a specific underlying file or directory. One common metaphor is that a directory is a folder that contains documents (files) and other folders (directories).

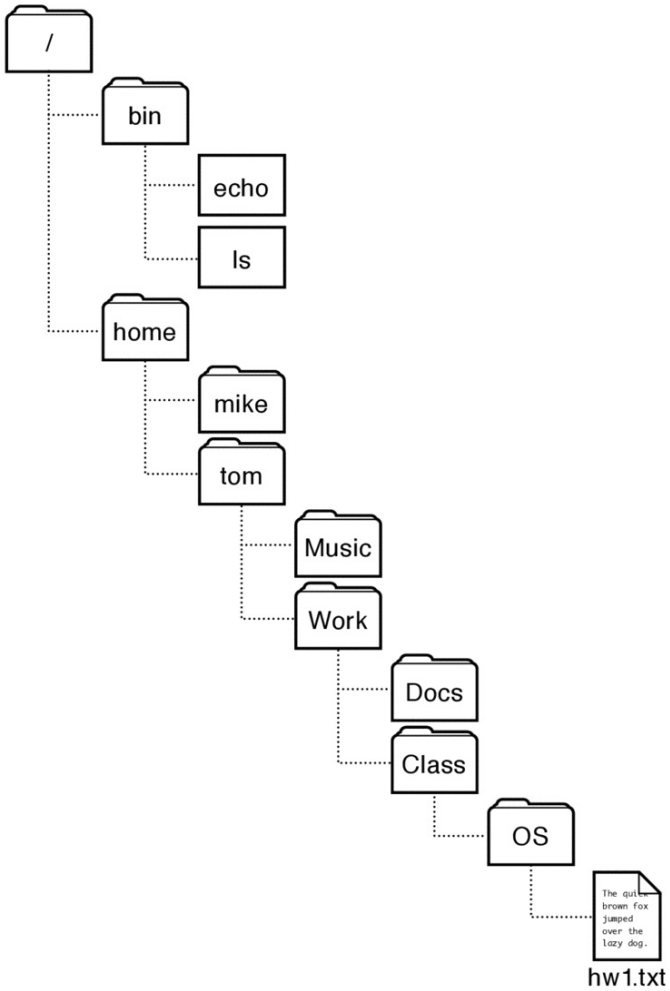


Figure 11.2: Example of a hierarchical organization of files using directories.

As Figure 11.2 illustrates, because directories can include names of other directories, they can be organized in a hierarchy so that different sets of associated files can be grouped in different directories. So, the directory /bin may include binary applications for your machine while /home/tom (Tom's "home directory") might include Tom's files. If Tom has many files, Tom's home directory may include additional directories to group them (e.g., /home/tom/Music and /home/tom/Work.) Each of these directories may have subdirectories (e.g.,/home/tom/Work/Class and /home/tom/ Work/Docs) and so on.

The string that identifies a file or directory (e.g., /home/tom/Work/Class/OS/hw1.txt or /home/tom) is called a *path*. Here, the symbol / (pronounced *slash*) separates components of the path, and each component represents an entry in a directory. So, hw1.txt is a file in the directory OS; OS is a directory in the directory Work; and so on.

If you think of the directory as a tree, then the root of the tree is a directory called, naturally enough, the *root directory*. Path names such as /bin/ls that begin with / define *absolute paths* that are interpreted relative to the root directory. So, /home refers to the directory called home in the root directory.

Path names such as Work/Class/OS that do not begin with / define *relative paths* that are interpreted by the operating system relative to a process's *current working directory*. So, if a process's current working directory is /home/tom, then the relative path Work/Class/OS is equivalent to the absolute path /home/tom/Work/Class/OS.

When you log in, your shell's current working directory is set to your *home directory*. Processes can change their current working directory with the chdir(path) system call. So, for example, if you log in and then type cd Work/Class/OS, your current working directory is changed from your home directory to the subdirectory Work/Class/OS in your home directory.

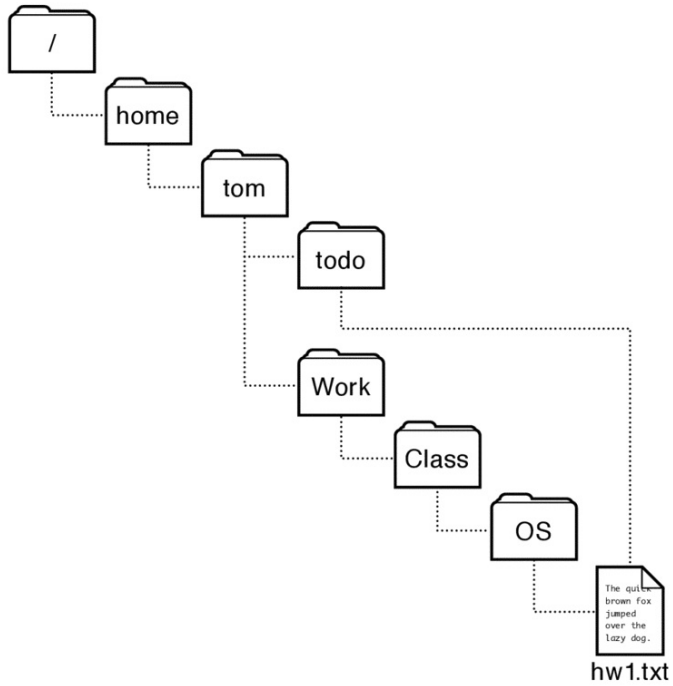


Figure 11.3: Example of a directed acyclic graph directory organization with multiple hard links to a file.

. and .. and ~

You may sometimes see path names in which directories are named ., .., or ~. For example,
> cd ~/Work/Class/OS
> cd ..
> ./a.out

., .., and ~ are special directory names in Unix. . refers to the *current directory*, .. refers to the *parent directory*, ~ refers to the *current user's home directory*, and *~name* refers to the *home directory of user name*.

So, the first shell command changes the current working directory to be the Work/ Class/OS directory in the user's home directory (e.g., /home/tom/Work/Class/OS). The second command changes the current working directory to be the Work/ Class directory in the user's home directory (e.g., ~/Work/Class or /home/ tom/Work/Class.) The third command executes the program a.out from the current working directory (e.g., ~/Work/Class/a.out or /home/tom/Work/Class/ a.out.)

If each file or directory is identified by exactly one path, then the directory hierarchy forms a tree. Occasionally, it is useful to have several different names for the same file or directory. For example, if you are actively working on a project, you might find it convenient to have the project appear in both your "todo" directory and a more permanent location (e.g., /home/tom/todo/hw1.txt and /home/tom/Work/Class/OS/hw1.txt as illustrated in Figure 11.3.)

The mapping between a name and the underlying file is called a *hard link*. If a system system allows multiple hard links to the same file, then the directory hierarchy may no longer be a tree. Most file systems that allow multiple hard links to a file restrict these links to avoid cycles, ensuring that their directory structures form a directed acyclic graph (DAG.) Avoiding cycles can simplify management by, for example, ensuring that recursive traversals of a directory structure terminate or by making it straightforward to use reference counting to garbage collect a file when the last link to it is removed.

In addition to hard links, many systems provide other ways to use multiple names to refer to the same file. See the sidebar for a comparison of hard links, soft links, symbolic links, shortcuts, and aliases.

Hard links, soft links, symbolic links, shortcuts, and aliases

A hard link is a directory mapping from a file name directly to an underlying file. As we will see in Chapter 13, directories will be implemented by storing mappings from *file names* to *file numbers* that uniquely identify each file. When you first create a file (e.g., /a/b), the directory entry you create is a hard link the the new file. If you then use link() to add another hard link to the file (e.g., link("/a/b", "/c/d"),) then both names are equally valid, independent names for the same underlying file. You could, for example, unlink("/a/b"), and /c/d would remain a valid name for the file.

Many systems also support *symbolic links* also known as *soft links*. A symbolic link is a directory mappings from a file name to *another file name*. If a file is opened via a symbolic link, the file system first translates the name in the symbolic link to the target name and then uses the target name to open the file. So, if you create /a/b , create a symbolic link from /c/d/ to /a/b, and then unlink /a/b, the file is no longer accessible and open("/c/d") will fail.

Although the potential for such dangling links is a disadvantage, symbolic links have a number of advantages over hard links. First, systems usually allow symbolic links to directories, not just regular files. Second, a symbolic link can refer to a file stored in a different file system or volume.

Some operating systems such as Microsoft Windows also support *shortcuts*, which appear similar to symbolic links but which are interpreted by the windowing system rather than by the file system. From the file system's point of view, a shortcut is just a regular file. The windowing system, however, treats shortcut files specially: when the shortcut file is selected via the windowing system, the windowing system opens that file, identifies the target file referenced by the shortcut, and acts as if the target file had been selected.

A MacOS file *alias* is similar to a symbolic link but with an added feature: if the target file is moved to have a new path name, the alias can still be used to reference the file.

Volume. Each instance of a file system manages files and directories for a volume. A *volume* is a collection of physical storage resources that form a logical storage device.

A volume is an abstraction that corresponds to a logical disk. In the simplest case, a volume corresponds to a single physical disk drive. Alternatively, a single physical disk can be partitioned and store multiple volumes or several physical disks can be combined so that a single volume spans multiple physical disks.

A single computer can make use of multiple file systems stored on multiple volumes by mounting multiple volumes in a single logical hierarchy. *Mounting* a volume on an existing file system creates a mapping from some path in the existing file system to the root directory of the mounted volume's file system and lets the mounted file system control mappings for all extensions of that path.

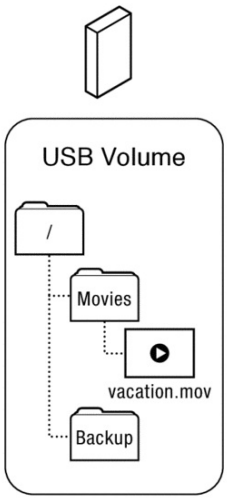


Figure 11.4: This USB disk holds a volume that is the physical storage for a file system.

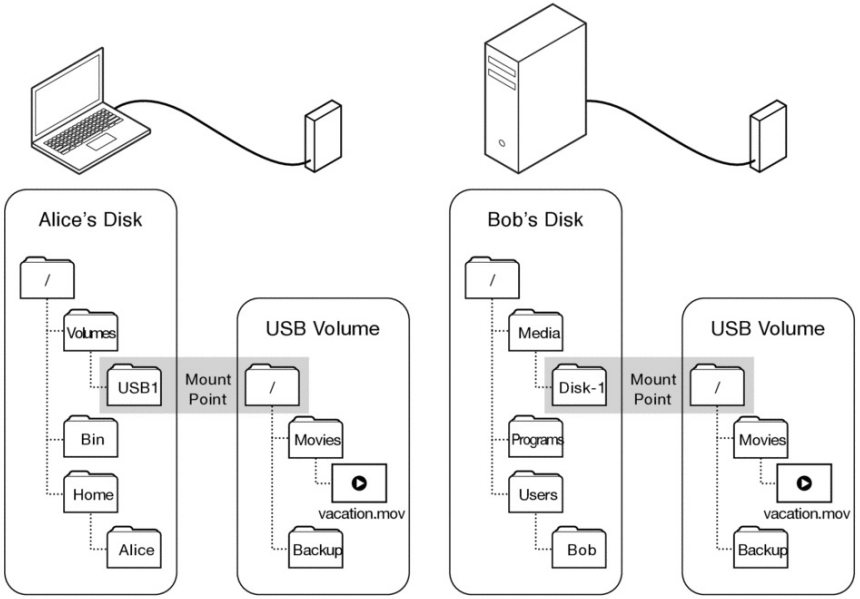


Figure 11.5: A volume can be mounted to another file system to join their directory hierarchies. For example, when the USB drive is connected to Alice's computer, she can access the vacation.mov movie using the path /Volumes/usb1/Movies/vacation.mov, and when the drive is connected to Bob's computer, he can access the movie using the path /media/disk-1/Movies/vacation.mov.

For example, suppose a USB drive contains a file system with the directories /Movies and /Backup as shown in Figure 11.4. If Alice plugs that drive into her laptop, the laptop's operating system might mount the USB volume's file system with the path /Volumes/usb1/ as shown in Figure 11.5. Then, if Alice calls open("/Volumes/usb1/ Movies/vacation.mov"), she will open the file /Movies/vacation.mov from the file system on the USB drive's volume. If, instead, Bob plugs that drive into his laptop, the laptop's operating system might mount the volume's file system with the path /media/disk-1, and Bob would access the same file using the path /media/disk-1/Movies/ vacation.mov.

11.2 API

Creating and deleting files

create (pathName)	Create a new file with the specified name.
link (existingName, newName)	Create a hard link — a new path name that refers to the same underlying file as an existing path name.
unlink (pathName)	Remove the specified name for a file from its directory; if that was the only name for the underlying file, then remove the file and free its resources.
mkdir (pathName)	Create a new directory with the specified name.
rmdir (pathName)	Remove the directory with the specified name.

Open and close

fileDescriptor open (pathName)	Prepare to access to the specified file (e.g., check access permissions and initialize kernel data structures for tracking per-process state of open files).
close (fileDescriptor)	Release resources associated with the specified open file.

File access

read (fileDescriptor, buf, len)	Read len bytes from the process's current position in the open file fileDescriptor and copy the results to a buffer buf in the application's memory.
write (fileDescriptor, len, buf)	Write len bytes of data from a buffer buf in the process's memory to the process's current position in the open file fileDescriptor.
seek (fileDescriptor, offset)	Change the process's current position in the open file fileDescriptor to the specified offset.
dataPtr mmap (fileDescriptor, off, len)	Set up a mapping between the data in the file fileDescriptor from off to off + len and an area in the application's virtual memory from dataPtr to dataPtr + len.
munmap (dataPtr, len)	Remove the mapping between the application's virtual memory and a mapped file.
fsync (fileDescriptor)	Force to disk all buffered, dirty pages for the file associated with fileDescriptor.

Figure 11.6: A simple API for accessing files.

For concreteness, Figure 11.6 shows a simple file system API for accessing files and directories.

Creating and deleting files. Processes create and destroy files with create() and unlink(). Create() does two things: it creates a new file that has initial metadata but no other data, and it creates a name for that file in a directory.

Link() creates a hard link — a new path name for an existing file. After a successful call to link(), there are multiple path names that refer to the same underlying file.

Unlink() removes a name for a file from its directory. If a file has multiple names or links, unlink() only removes the specified name, leaving the file accessible via other names. If the specified name is the last (or only) link to a file, then unlink() also deletes the underlying file and frees its resources.

Mkdir() and rmdir() create and delete directories.

EXAMPLE: Linking to files vs. linking to directories. Systems such as Linux support a link() system call, but they do not allow new hard links to be created to a directory. E.g., existingPath must not be a directory. Why does Linux mandate this restriction?

ANSWER: Preventing multiple hard links to a directory prevents cycles, ensuring that the directory structure is always a directed acyclic graph (DAG).

Additionally, allowing hard links to a directory would muddle a directory's parent directory entry (e.g., ".." as discussed in the sidebar). □

Open and close. To start accessing a file, a process calls open() to get a *file descriptor* it can use to refer to the open file. *File descriptor* is Unix terminology; in other systems the descriptor may be called a *file handle* or a *file stream*.

Operating systems require processes to explicitly open() files and access them via file descriptors rather than simply passing the path name to read() and write() calls for two reasons. First, path parsing and permission checking can be done just when a file is opened and need not be repeated on each read or write. Second, when a process opens a file, the operating system creates a data structure that stores information about the process's open file such as the file's ID, whether the process can write or just read the file, and a pointer to the process's current position within the file. The file descriptor can thus be thought of as a reference to the operating system's per-open-file data structure that the operating system will use for managing the process's access to the file.

When an application is done using a file, it calls close(), which releases the open file record in the operating system.

File access. While a file is open, an application can access the file's data in two ways. First, it can use the traditional procedural interface, making system calls to read() and write() on an open file. Calls to read() and write() start from the process's current file position, and they advance the current file position by the number of bytes successfully read or written. So, a sequence of read() or write() calls moves sequentially through a file. To support random access within a file, the seek() call changes a process's current position for a specified open file.

Rather than using read() and write() to access a file's data, an application can use mmap() to establish a mapping between a region of the process's virtual memory and some region of the file. Once a file has been mapped, memory loads and stores to that virtual memory region will read and write the file's data either by accessing a shared page from the kernel's file cache, or by triggering a page fault exception that causes the kernel to fetch the desired page of data from the file system into memory. When an application is done with a file, it can call munmap() to remove the mappings.

Finally, the fsync() call is important for reliability. When an application updates a file via a write() or a memory store to a mapped file, the updates are buffered in memory and written back to stable storage at some future time. Fsync() ensures that all pending updates for a file are written to persistent storage before the call returns. Applications use this function for two purposes. First, calling fsync() ensures that updates are durable and will not be lost if there is a crash or power failure. Second, calling fsync() between two updates ensures that the first is written to persistent storage before the second. Note that calling fsync() is not always necessary; the operating system ensures that all updates are made durable by periodically flushing all dirty file blocks to stable storage.

Modern file access APIs

The API shown in Figure 11.6 is similar to most widely used file access APIs, but it is somewhat simplified.

For example, each of the listed calls is similar to a call provided by the POSIX interface, but the API shown in Figure 11.6 omits some arguments and options found in POSIX. The POSIX open() call, for example, includes two additional arguments one to specify various flags such as whether the file should be opened in read-only or read-write mode and the other to specify the access control permissions that should be used if the open() call creates a new file.

In addition, real-world file access APIs are likely to have a number of additional calls. For example, the Microsoft Windows file access API includes dozens of calls including calls to lock and unlock a file, to encrypt and decrypt a file, or to find a file in a directory whose name matches a specific pattern.

11.3 Software Layers

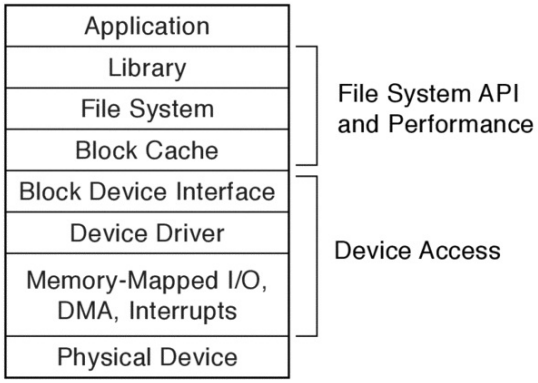


Figure 11.7: Layered abstractions provide access to I/O systems such as storage systems.

As shown in Figure 11.7, operating systems implement the file system abstraction through a series of software layers. Broadly speaking, these layers have two sets of tasks:

- **API and performance.** The top levels of the software stack — user-level libraries, kernel-level file systems, and the kernel's block cache — provide a convenient API for accessing named files and also work to minimize slow storage accesses via caching, write buffering, and prefetching.
- **Device access.** Lower levels of the software stack provide ways for the operating system to access a wide range of I/O devices. Device drivers hide the details of specific I/O hardware by providing hardware-specific code for each device, and placing that code behind a simpler, more general interfaces that the rest of the operating system can use such as a block device interface. The device drivers execute as normal kernel-level code, using the systems' main processors and memory, but they must interact with the I/O devices. A system's processors and memory communicate with its I/O devices using Memory-Mapped I/O, DMA, and Interrupts.

In the rest of this section, we first talk about the file system API and performance layers. We then discuss device access.

11.3.1 API and Performance

The top levels of the file system software stack — divided between application libraries and operating system kernel code — provide the file system API and also provide caching and write buffering to improve performance.

System calls and libraries. The file system abstraction such as the API shown in Figure 11.6 can be provided directly by system calls. Alternatively, application libraries can wrap the system calls to add additional functionality such as buffering.

For example, in Linux, applications can access files directly using system calls (e.g., open(), read(), write(), and close().) Alternatively, applications can use the stdio library calls (e.g., fopen(), fread(), fwrite(), and fclose()). The advantage of the latter is that the library includes buffers to aggregate a program's small reads and writes into system calls that access larger blocks, which can reduce overheads. For example, if a program uses the library function fread() to read 1 byte of data, the fread() implementation may use the read() system call to read a larger block of data (e.g., 4 KB) into a buffer maintained by the library in the application's address space. Then, if the process calls fread() again to read another byte, the library just returns the byte from the buffer without needing to do a system call.

Block cache. Typical storage devices are much slower than a computer's main memory. The operating system's block cache therefore caches recently read blocks, and it buffers recently written blocks so that they can be written back to the storage device at a later time.

In addition to improving performance by caching and write buffering, the block cache serves as a synchronization point: because all requests for a given block go through the block cache, the operating system includes information with each buffer cache entry to, for example, prevent one process from reading a block while another process writes it or to ensure that a given block is only fetched from the storage device once, even if it is simultaneously read by many processes.

Prefetching. Operating systems use prefetching to improve I/O performance. For example, if a process reads the first two blocks of a file, the operating system may prefetch the next ten blocks.

Such prefetching can have several beneficial effects:

- **Reduced latency.** When predictions are accurate, prefetching can help the latency of future requests because reads can be serviced from main memory rather than from slower storage devices.
- **Reduced device overhead.** Prefetching can help reduce storage device overheads by replacing a large number of small requests with one large one.
- **Improved parallelism.** Some storage devices such as Redundant Arrays of Inexpensive Disks (RAIDs) and Flash drives are able to process multiple requests at once, in parallel. Prefetching provides a way for operating systems to take advantage of available hardware parallelism.

Prefetching, however, must be used with care. Too-aggressive prefetching can cause problems:

- **Cache pressure.** Each prefetched block is stored in the block cache, and it may displace another block from the cache. If the evicted block is needed before the prefetched one is used, prefetching is likely to hurt overall performance.
- **I/O contention.** Prefetch requests consume I/O resources. If other requests have to wait behind prefetch requests, prefetching may hurt overall performance.
- **Wasted effort.** Prefetching is speculative. If the prefetched blocks end up being needed, then prefetching can help performance; otherwise, prefetching may hurt overall performance by wasting memory space, I/O device bandwidth, and CPU cycles.

11.3.2 Device Drivers: Common Abstractions

Device drivers translate between the high level abstractions implemented by the operating system and the hardware-specific details of I/O devices.

An operating system may have to deal with many different I/O devices. For example, a laptop on a desk might be connected to two keyboards (one internal and one external), a trackpad, a mouse, a wired ethernet, a wireless 802.11 network, a wireless bluetooth network, two disk drives (one internal and one external), a microphone, a speaker, a camera, a printer, a scanner, and a USB thumb drive. And that is just a handful of the literally thousands of devices that could be attached to a computer today. Building an operating system that treats each case separately would be impossibly complex.

Layering helps simplify operating systems by providing common ways to access various classes of devices. For example, for any given operating system, storage device drivers typically implement a standard *block device* interface that allows data to be read or written in fixed-sized blocks (e.g., 512, 2048, or 4096 bytes).

Such a standard interface lets an operating system easily use a wide range of similar devices. A file system implemented to run on top of the standard block device interface can store files on any storage device whose driver implements that interface, be it a Seagate spinning disk drive, an Intel solid state drive, a Western Digital RAID, or an Amazon Elastic Block Store volume. These devices all have different internal organizations and control registers, but if each manufacturer provides a device driver that exports the standard interface, the rest of the operating system does not need to be concerned with these per-device details.

Challenge: device driver reliability

Because device drivers are hardware-specific, they are often written and updated by the hardware manufacturer rather than the operating system's main authors. Furthermore, because there are large numbers of devices — some operating systems support tens of thousands of devices — device driver code may represent a large fraction of an operating system's code.

Unfortunately, bugs in device drivers have the potential to affect more than the device. A device driver usually runs as part of the operating system kernel since kernel routines depend on it and because it needs to access the hardware of its device. However, if the device driver is part of the kernel, then a device driver's bugs have the potential to affect the overall reliability of a system. For example, in 2003 it was reported that drivers caused about 85% of failures in the Windows XP operating system.

To improve reliability, operating systems are increasingly using protection techniques similar to those used to isolate user-level programs to isolate device drivers from the kernel and from each other.

11.3.3 Device Access

How should an operating system's device drivers communicate with and control a storage device? At first blush, a storage device seems very different from the memory and CPU resources we have discussed so far. For example, a disk drive includes several motors, a sensor for reading data, and an electromagnet for writing data.

Memory-mapped I/O. As Figure 11.8 illustrates, I/O devices are typically connected to an I/O bus that is connected to the system's memory bus. Each I/O device has controller with a set of registers that can be written and read to transmit commands and data to and from the device. For example, a simple keyboard controller might have one register that can be read to learn the most recent key pressed and another register than can be written to turn the caps-lock light on or off.

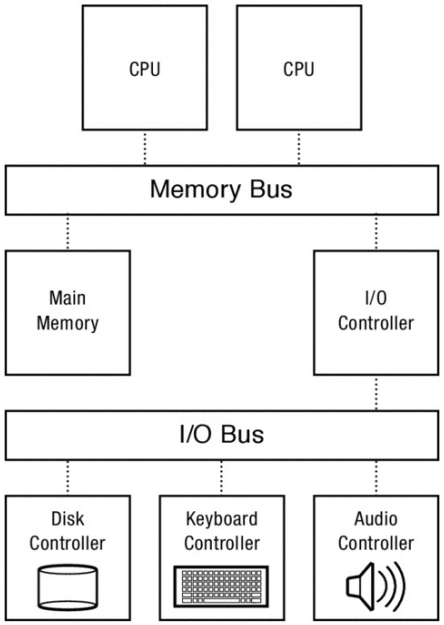


Figure 11.8: I/O devices are attached to the I/O bus, which is attached to the memory bus.

To allow I/O control registers to be read and written, systems implement memory-mapped I/O. *Memory-mapped I/O* maps each device's control registers to a range of physical addresses on the memory bus. Reads and writes by the CPU to this physical address range do not go to main memory. Instead, they go to registers on the I/O devices's controllers. Thus, the operating system's keyboard device driver might learn the value of the last key pressed by reading from physical address, say, 0xC00002000.

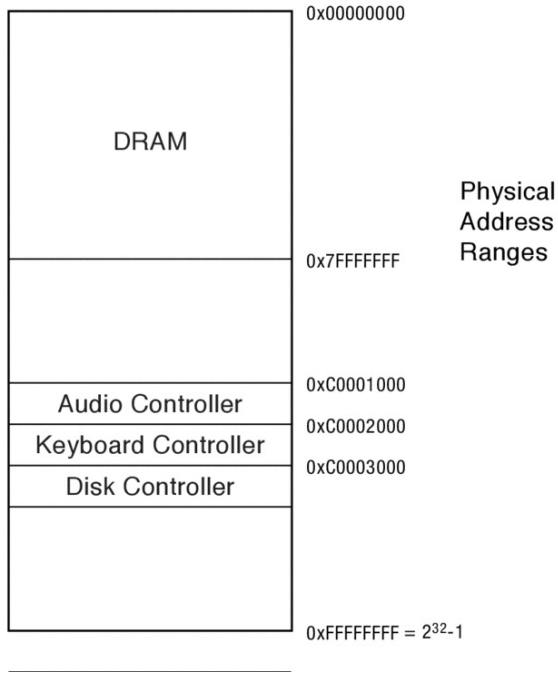


Figure 11.9: Physical address map for a system with 2 GB of DRAM and 3 memory-mapped I/O devices.

The hardware maps different devices to different physical address ranges. Figure 11.9 shows the physical address map for a hypothetical system with a 32 bit physical address space capable of addressing 4 GB of physical memory. This system has 2 GB of DRAM in it, consuming physical addresses 0x00000000 (0) to 0x7FFFFFFF (2^{31} - 1). Controllers for each of its three I/O devices are mapped to ranges of addresses in the first few kilobytes above 3 GB. For example, physical addresses from 0xC0001000 to 0xC0001FFF access registers in the disk controller.

Port mapped I/O

Today, memory-mapped I/O is the dominant paradigm for accessing I/O device's control registers. However an older style, *port mapped I/O*, is still sometimes used. Notably, the x86 architecture supports both memory-mapped I/O and port mapped I/O.

Port mapped I/O is similar to memory-mapped I/O in that instructions read from and write to specified addresses to control I/O devices. There are two differences. First, where memory-mapped I/O uses standard memory-access instructions (e.g., load and store) to communicate with devices, port mapped I/O uses distinct I/O instructions. For example, the x86 architecture uses the in and out instructions for port mapped I/O. Second, whereas memory-mapped I/O uses the same physical address space as is used for the system's

main memory, the address space for port mapped I/O is distinct from the main memory address space.

For example, in x86 I/O can be done using either memory-mapped or port mapped I/O, and the low-level assembly code is similar for both cases:

Memory mapped I/O

```
MOV register, memAddr // To read
MOV memAddr, register // To write
```

Port mapped I/O

```
IN register, portAddr  // To read
OUT portAddr, register // To write
```

Port mapped I/O can be useful in architectures with constrained physical memory addresses since I/O devices do not need to consume ranges of physical memory addresses. On the other hand, for systems with sufficiently large physical address spaces, memory-mapped I/O can be simpler since no new instructions or address ranges need to be defined and since device drivers can use any standard memory access instructions to access devices. Also, memory-mapped I/O provides a more unified model for supporting DMA — direct transfers between I/O devices and main memory.

DMA. Many I/O devices, including most storage devices, transfer data in bulk. For example, operating systems don't read a word or two from disk, they usually do transfers of at least a few kilobytes at a time. Rather than requiring the CPU to read or write each word of a large transfer, I/O devices can use direct memory access. When using *direct memory access (DMA)*, the I/O device copies a block of data between its own internal memory and the system's main memory.

To set up a DMA transfer, a simple operating system could use memory-mapped I/O to provide a target physical address, transfer length, and operation code to the device. Then, the device copies data to or from the target address without requiring additional processor involvement.

After setting up a DMA transfer, the operating system must not use the target physical pages for any other purpose until the DMA transfer is done. The operating system therefore "pins" the target pages in memory so that they cannot be reused until they are unpinned. For example, a pinned physical page cannot be swapped out to disk and then remapped to some other virtual address.

Advanced DMA

Although a setting up a device's DMA can be as simple as providing a target physical address and length and then saying "go!", more sophisticated interfaces are increasingly used.

For example rather than giving devices direct access to the machine's physical address space, some systems include an I/O memory management unit (IOMMU) that translates device virtual addresses to main memory physical addresses similar to how a processor's TLB translates processor virtual addresses to main memory physical addresses. An IOMMU can provide both protection (e.g., preventing a buggy IO device from overwriting arbitrary memory) and simpler abstractions (e.g., allowing devices to use virtual addresses so that, for example, a long transfer can be made to a range of consecutive virtual pages rather than a collection of physical pages scattered across memory.)

Also, some devices add a level of indirection so that they can interrupt the CPU less often. For example, rather than using memory mapped I/O to set up each DMA request, the CPU and I/O device could share two lists in memory: one list of pending I/O requests and another of completed I/O requests. Then, the CPU could set up dozens of disk requests and only be interrupted when all of them are done.

Sophisticated I/O devices can even be configured to take different actions depending the data they receive. For example, some high performance network interfaces parse incoming packets and direct interrupts to different processors based on the network connection to which a received packet belongs.

Interrupts. The operating system needs to know when I/O devices have completed handling a request or when new external input arrives. One option is *polling*, repeatedly using memory-mapped I/O to read a status register on the device. Because I/O devices are often much slower than CPUs and because inputs received by I/O devices may arrive at irregular rates, it us usually better for I/O devices to use an *interrupt* to notify the operating system of important events.

11.3.4 Putting It All Together: A Simple Disk Request

When a process issues a system call like read() to read data from disk into the process's memory, the operating system moves the calling thread to a wait queue. Then, the operating system uses memory-mapped I/O both to tell the disk to read the requested data and to set up DMA so that the disk can place that data in the kernel's memory. The disk then reads the data and DMAs it into main memory; once that is done, the disk triggers an interrupt. The operating system's interrupt handler then copies the data from the kernel's buffer into the process's address space. Finally, the operating system moves the thread the ready list. When the thread next runs, it will returns from the system call with the data now present in the specified buffer.

11.4 Summary and Future Directions

The file system interface is a stable one, and small variations of interface described here can be found in many operating systems and for many storage devices.

Yet, the file system abstraction is imperfect, and application writers need to use it carefully to get acceptable performance and reliability. For example, if an application write()s a file, the update may not be durable when the write() call returns; application writers often call fsync() to ensure durability of data.

Could better file system APIs simplify programming? For example, if file systems allowed users to update multiple objects atomically, that might simplify many applications that currently must carefully constrain the order that their updates are stored using crude techniques such as using fsync as a barrier between one set of updates and the next.

Could better file system APIs improve performance? For example, one proposed interface allows an application to direct the operating system to transfer a range of bytes from a file to a network connection. Such an interface might, for example, reduce overheads for a movie server that streams movies across a network to clients.

Exercises

1. **Discussion** Suppose a process successfully opens an existing file that has a single hard link to it, but while the process is reading that file, another process unlinks that file. What should happen to subsequent reads by the first process? Should they succeed? Should they fail? Why?
2. In Linux, suppose a process successfully opens an existing file that has a single hard link to it, but while the process is reading that file, another process unlinks that file? What happens to subsequent reads by the first process? Do they succeed? Do they fail? (Answer this problem by consulting documentation or by writing a program to test the behavior of the system in this case.)
3. Write a program that creates a new file, writes 100KB to it, flushes the writes, and deletes it. Time how long each of these steps takes.

   Hint You may find the POSIX system calls creat(), write(), fflush(), close(), and gettimeofday() useful. See the manual pages for details on how to use these.
4. Consider a text editor that saves a file whenever you click a save button. Suppose that when you press the button, the editor simply (1) animates the button “down” event (e.g., by coloring the button grey), (2) uses the write() system call to write your text to your file, and then (3) animates the button “up” event (e.g., by coloring the button white). What bad thing could happen if a user edits a file, saves it, and then turns off her machine by flipping the power switch (rather than shutting the machine down cleanly)?
5. Write a program that times how long it takes to issue 100,000 one-byte writes in each of two ways. First, time how long it takes to use the POSIX system calls creat(), write(), and close() directly. Then see how long these writes take if the program uses the stdio library calls (e.g., fopen(), fwrite(), and fclose()) instead. Explain your results.

12. Storage Devices

Treat disks like tape. —*John Ousterhout*

Although today's persistent storage devices have large capacity and low cost, they have drastically worse performance than volatile DRAM memory.

Not only that, but the characteristics are different and are peculiar to specific persistent storage devices. For example, although programs can access random individual words of DRAM with good performance, programs can only access today's disk and flash storage devices hundreds or thousands of bytes at a time. Furthermore, even if an application restricts itself to supported access sizes (e.g., 2 KB per read or write), if the application access pattern is random, the application may be slower by a factor of several hundred than if the application accessed the same amount of data sequentially.

To cope with the limitations and to maximize the performance of storage devices, both file system designers and application writers need to understand the physical characteristics of persistent storage devices.

Chapter roadmap. This chapter discusses two types of persistent storage: *magnetic disks* and *flash memory*. Both are widely used. Magnetic disks provide persistent storage for most servers, workstations, and laptops. Flash memory provides persistent storage for most smart phones, tablets, and cameras and for an increasing fraction of laptops.

12.1 Magnetic Disk

Figure 12.1: A partially-disassembled magnetic disk drive.

Magnetic disk is a non-volatile storage technology that is widely used in laptops, desktops, and servers. Disk drives work by magnetically storing data on a thin metallic film bonded to a glass, ceramic, or aluminum disk that rotates rapidly. Figure 12.1 shows a disk drive without its protective cover, and Figure 12.2 shows a schematic of a disk drive, identifying key components.

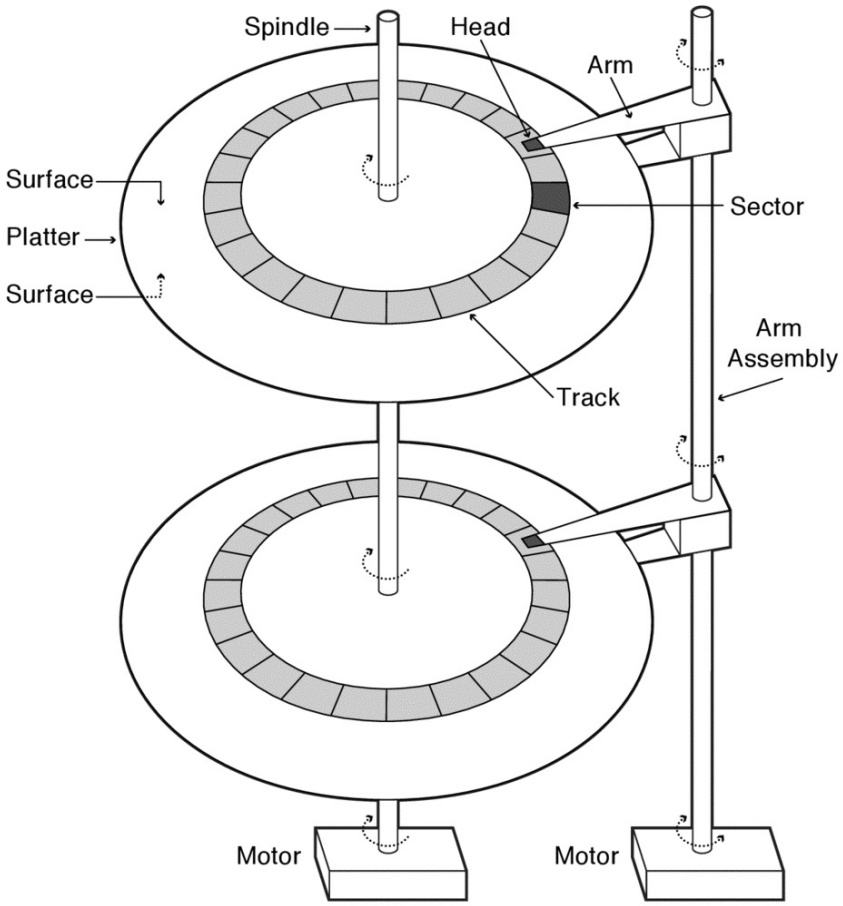


Figure 12.2: Key components of a magnetic disk drive.

Each drive holds one or more *platters*, thin round plates that hold the magnetic material. Each platter has two *surfaces*, one on each side. When the drive powers up, the platters are constantly spinning on a *spindle* powered by a *motor*. In 2011, disks commonly spin at 4200–15000 RPM (70–250 revolutions per second.)

A disk *head* is the component that reads and writes data by sensing or introducing a magnetic field on a surface. There is one head per surface, and as a surface spins underneath a head, the head reads or writes a sequence of bits along a circle centered on the disk's spindle. As a disk platters spins, it creates a layer of rapidly spinning air, and the disk head floats on that layer, allowing the head to get extremely close to the platter without contacting it. A *head crash* occurs when the disk head breaks through this layer with enough force to damage the magnetic surface below; head crashes can be caused by excessive shock such as dropping a running drive.

To reach the full surface, each disk head attaches to an *arm*, and all of a disk's arms attach to a single *arm assembly* that includes a motor that can move the arms across the surfaces of the platters. Note that an assembly has just one motor, and all of its arms move together.

Data bits are stored in fixed-size *sectors*; typically, sectors are 512 bytes. The disk hardware cannot read or write individual bytes or words; instead, it must always read or write at least an entire sector. This means that to change one byte in a sector, the operating system must read the old sector, update the byte in memory, and rewrite the entire sector to disk. One reason for this restriction is that the disk encodes each sector with additional error correction code data, allowing it to fix (or at least detect) imperfectly read or written data, which, in turn allows higher density storage and higher bandwidth operation.

A circle of sectors on a surface is called a *track*. The disk can read or write all of the data on a track without having to move the disk arm, so reading or writing a sequence of sectors on the same track is much faster than reading or writing sectors on different tracks.

To maximize sequential access speed, logical sector zero on each track is staggered from sector zero on the previous track by an amount corresponding to time it takes the disk to move the head from one track to another or to switch from the head on one surface to the head on another one. This staggering is called *track skewing*.

To increase storage density and disk capacity, disk manufacturers make tracks and sectors as thin and small as possible. If there are imperfections in a sector, then that sector may be unable to reliably store data. Manufacturers therefore include spare sectors distributed across each surface. The disk firmware or the file system's low-level formatting can then use *sector sparing* to remap sectors to use spare sectors instead of faulty sectors. *Slip sparing* helps retain good sequential access performance by remapping all sectors from the bad sector to the next spare, advancing each logical sector in that range by one physical sector on disk.

Disk drives often include a few MB of *buffer memory*, memory that the disk's controller uses to buffer data being read from or written to the disk, for track buffering, and for write acceleration.

Track buffering improves performance by storing sectors that have been read by the disk head but have not yet been requested by the operating system. In particular, when a disk head moves to a track, it may have to wait for the sector it needs to access to rotate under the disk head. While the disk is waiting, it reads unrequested sectors to its rack buffer so that if the operating system requests those sectors later, they can be returned immediately.

Write acceleration stores data to be written to disk in the disk's buffer memory and acknowledges the writes to the operating system before the data is actually written to the platter; the disk firmware flushes the writes from the track buffer to the platter at some later time. This technique can significantly increase the apparent speed of the disk, but it carries risks — if power is lost before buffered data is safely stored, then data might be lost.

Server drives implementing the SCSI or Fibre Channel interfaces and increasing numbers of commodity drives with the Serial ATA (SATA) interface implement a safer form of write acceleration with *tagged command queueing* (TCQ) (also called *native command queueing* (NCQ) for SATA drives.) TCQ allows an operating system to issue multiple concurrent requests to the disk and for the disk to process those requests out of order to optimize scheduling, and it can be configured to only acknowledge write requests when the blocks are safely on the platter.

12.1.1 Disk Access and Performance

Operating systems send commands to a disk to read or write one or more consecutive sectors. A disk's sectors are identified with *logical block addresses* (LBAs) that specify the surface, track, and sector to be accessed.

To service a request for a sequence of blocks starting at some sector, the disk must first seek to the right track, wait for the first desired sector to rotate to the head, and then transfer the blocks. Therefore, the time for a disk access is:

disk access time = seek time + rotation time + transfer time

- **Seek.** The disk must first *seek* — move its arm over the desired track. To seek, the disk first activates a motor that moves the arm assembly to approximately the right place on disk. Then, as arm stops vibrating from the motion of the seek, the disk begins reading positioning information embedded in the sectors to determine exactly where it is and to make fine-grained positioning corrections to *settle* on the desired track. Once the head has settled on the right track, the disk uses signal strength and positioning information to make minute corrections in the arm position to keep the head over the desired track.

  A request's seek time depends on how far the disk arm has to move.

  A disk's *minimum seek time* is the time it takes for the head to move from one track to an adjacent one. For short seeks, disks typically just "resettle" the head on the new track by updating the target track number in the track-following circuitry. Minimum seek times of 0.3–1.5 ms are typical.

  If a disk is reading the tth track on one surface, its *head switch time* is the time it would take to begin reading the tth track on a different surface. Tracks can be less than a micron wide and tracks on different surfaces are not perfectly aligned. So, a head switch between the same tracks on different surfaces has a cost similar to a minimum seek: the disk begins using the sensor on a different head and then resettles the disk on the desired track for that surface.

  A disk's *maximum seek time* is the time it takes the head to move from the innermost track to the outermost one or vice versa. Maximum seek times are typically over 10 ms and can be over 20 ms.

  A disk's *average seek time* is the average across seeks between each possible pair of tracks on a disk. This value is often approximated as the time to seek one third of the way across the disk.

Beware of "average seek time"

Although the name *average seek time* makes it tempting to use this metric when estimating the time it will take a disk to complete a particular workload, it is often the wrong metric to use. Average seek time — the average across seeks between each possible pair of tracks on disk — was defined this way to make it a well-defined, standard metric, not because it is representative of common workloads.

The definition of average seek time essentially assumes no locality in a workload, so it is very nearly a worst-case scenario. Many workloads access sectors that are likely to be near one another; for example, most operating systems attempt to place files sequentially on disk and to place different files in a directory on the same track or on tracks near one another. For these (common) workloads, the seek times observed may be closer to the disk's minimum seek time than its "average" seek time.

The demise of the cylinder

A *cylinder* on a disk is a set of tracks on different surfaces with the same track index. For example, on a 2-platter drive, the 8th tracks on surfaces 0, 1, 2, and 3 would form the 8th cylinder of the drive.

Some early file systems put related data on different surfaces but in the same cylinder. The idea was that data from the different tracks in the cylinder could be read without a requiring a seek. Once a cylinder was full, the file system would start placing data in one of the adjacent cylinders.

As disk densities have increased, the importance of the cylinder has declined. Today, a disk's tracks can be less than a micron wide. To follow a track at these densities, a controller monitors the signals from a disk's head to control the disk arm assembly's motor to keep the head centered on a track. Furthermore, at these densities, the tracks of a cylinder may not be perfectly aligned. As a result, when a disk switches disk heads, the new head must center itself over the desired track. So, switching heads within a cylinder ends up being similar to a short 1-track seek: the controller chooses the new cylinder/track and the disk head settles over the target track. Today, accessing different tracks within the same cylinder costs about the same as accessing adjoining tracks on the same platter.

- **Rotate.** Once the disk head has settled on the right track, it must wait for the target sector to rotate under it. This waiting time is called the *rotational latency*. Today, most disks rotate at 4200 RPM to 15,000 RPM (15 ms to 4 ms per rotation), and for

many workloads a reasonable estimate of rotational latency is one-half the time for a full rotation — 7.5 ms–2 ms.

Once a disk head has settled on a new track, most disks immediately begin reading sectors into their buffer memory, regardless of which sectors have been requested. This way, if there is a request for one of the sectors that have already passed under the disk head, the data can be transferred immediately, rather than having to delay the request for nearly a full rotation to reread the data.

- **Transfer.** Once the disk head reaches a desired sector, the disk must transfer the data from the sector to its buffer memory (for reads) or vice versa (for writes) as the sectors rotate underneath the head. Then, for reads, it must transfer the data from its buffer memory to the host's main memory. For writes, the order of the transfers is reversed.

  To amortize seek and rotation time, disk requests are often for multiple sequential sectors. The time to transfer one or more sequential sectors from (or to) a surface once the disk head begins reading (or writing) the first sector is the *surface transfer time*.

  On a modern disk, the surface transfer time for a single sector is much smaller than the seek time or rotational latency. For example, disk bandwidths often exceed 100 MB/s, so the surface transfer time for a 512-byte sector is often under 5 microseconds (0.005 ms).

  Because a disk's outer tracks have room for more sectors than its inner tracks and because a given disk spins at a constant rate, the surface transfer bandwidth is often higher for the outer tracks than the inner tracks.

  For a disk read, once sectors have been transferred to the disk's buffer memory, they must be transferred to the host's memory over some connection such as SATA (serial ATA), SAS (serial attached SCSI), Fibre Channel, or USB (universal serial bus). For writes, the transfer goes in the other direction. The time to transfer data between the host's memory and the disk's buffer is the *host transfer time*. Typical bandwidths range from 60 MB/s for USB 2.0 to 2500 MB/s for Fibre Channel-20GFC.

  For multi-sector reads, disks pipeline transfers between the surface and disk buffer memory and between buffer memory and host memory; so for large transfers, the total transfer time will be dominated by whichever of these is the bottleneck. Similarly, for writes, disks overlap the host transfer with the seek, rotation, and surface transfer; again, the total transfer time will be dominated by whichever is the bottleneck.

12.1.2 Case Study: Toshiba MK3254GSY

Figure 12.3 shows some key parameters for a recent 2.5-inch disk drive for laptop computers.

Size	
Platters/Heads	2/4
Capacity	320 GB
Performance	
Spindle speed	7200 RPM
Average seek time read/write	10.5 ms / 12.0 ms
Maximum seek time	19 ms
Track-to-track seek time	1 ms
Transfer rate (surface to buffer)	54–128 MB/s
Transfer rate (buffer to host)	375 MB/s
Buffer memory	16 MB
Power	
Typical	16.35 W
Idle	11.68 W

Figure 12.3: Hardware specifications for a laptop disk (Toshiba MK3254GSY) manufactured in 2008.

This disk stores 320 GB of data on two platters, so it stores 80 GB per surface. The platters spin at 7200 revolutions per minute, which is 8.3 ms per revolution; since each platter's diameter is about 6.3 cm, the outer edge of each platter is moving at about 85 km/hour!

The disk's data sheet indicates an average seek time for the drive of 10.5 ms for reads and 12.0 ms for writes. The seek time for reads and writes differs because the disk starts attempting to read data before the disk arm has completely settled, but it must wait a bit longer before it is safe to write.

When transferring long runs of contiguous sectors, the disk's bandwidth is 54-128 MB/s. The bandwidth is expressed as a range because the disk's outer tracks have more sectors than its inner tracks, so when the disk is accessing data on its outer tracks, sectors sweep past the disk head at a higher rate.

Once the data is transferred off the platter, the disk can send it to the main memory of the

computer at up to 375 MB/s via a SATA (Serial ATA) interface.

Random vs. sequential performance. Given seek and rotational times measured in milliseconds, small accesses to random sectors on disk are much slower than large, sequential accesses.

EXAMPLE: Random access workload. For the disk described in Figure 12.3, consider a workload consisting of 500 read requests, each of a randomly chosen sector on disk, assuming requests are serviced in FIFO order. How long will servicing these requests take?

ANSWER: Disk access time is seek time + rotation time + transfer time.

Seek time. Each request requires a seek from a random starting track to a random ending track, so the disk's average seek time of 10.5 ms is a good estimate of the cost of each seek.

Rotation time. Once the disk head settles on the right track, it must wait for the desired sector to rotate under it. Since there is no reason to expect the desired sector to be particularly near or far from the disk head when it settles, a reasonable estimate for rotation time is 4.15 ms, one half of the time that it takes a 7200 RPM disk to rotate once.

Transfer time. The disk's surface bandwidth is at least 54 MB/s, so transferring 512 bytes takes at most 9.5 µS (0.0095 ms).

Total time. 10.5 + 4.15 + .0095 = 14.66 ms per request, so **500 requests will take about 7.33 seconds**. □

EXAMPLE: Sequential access workload. For the disk described in Figure 12.3, consider a workload consisting of a read request for 500 sequential sectors on the same track. How long will servicing these requests take?

ANSWER: Disk access time is seek time + rotation time + transfer time.

Seek time. Since we do not know which track we are starting with or which track we are reading from, we use the average seek time, 10.5 ms, as an estimate for the seek time.

Rotation time. Since we don't know the position of the disk when the request is issued, a simple and reasonable estimate for the time for the first desired block to rotate to the disk head is 4.15 ms, one half of the time that it takes a 7200 RPM disk to rotate once.

Transfer time. A simple estimate is that 500 sectors can be transferred in 4.8 to 2.0 ms, depending on whether they are on the inner or outer tracks.

500 sectors × 512 bytes/sector

× 1 / (54 ×10^6 bytes/second) = 4.8 ms

500 sectors × 512 bytes/sector

× 1 / (128 ×10^6 bytes/second) = 2 ms

(Too) simple answer. These three estimates give us a range from

10.5 + 4.15 + 2 = 16.7 ms

10.5 + 4.15 + 4.8 = 19.5 ms

More precise answer. However, this simple answer ignores the track buffer. Since the transfer time is a large fraction of the rotation time (about 1/4 to 1/2 of the time for a full rotation), we know that the request covers a significant fraction of a track. This means that there is a good chance that after the seek and settle time, the disk head will be in the middle of the region to be read. In this case, the disk will immediately read some of the track into the track buffer; then it will wait for the first track to rotate around; then it will read the remainder of the desired data.

We can estimate that for the outer track, there is a one in four chance that the initial seek and settle will finish while the head is within the desired range of sectors, and that when that happens, we read an average of 1⁄8th of the desired data before we arrive at the first desired sector. So, for the outer track, this overlap will save us 1⁄4 × 1⁄8 = 1⁄32 of a rotation for the average transfer. This effect slightly reduces the average access time: 16.7 ms - (1⁄32) × 8.3 ms = 16.4 ms.

Similarly, for the inner tracks, there is about a one in two chance that the initial seek will settle in the middle of the desired data, saving on average 1⁄2 × 1⁄4 = 1⁄8. This reduces the average access time: 19.5 ms - (1⁄8) × 8.3 ms = 18.5 ms.

So, we estimate that **such an access would take between 16.4 ms and 18.5 ms**. □

Notice that the sequential workload takes vastly less time than the random workload (less than 20 milliseconds vs. 5.5 seconds). This orders of magnitude disparity between sequential and random access performance influences many aspects of file system design and use.

Still, even for a 500 sector request, a non-trivial amount of the access time is spent seeking and rotating rather than transferring.

EXAMPLE: Effective bandwidth. For the transfer of 500 sequential sectors examined in the previous example, what fraction of the disk's surface bandwidth is realized?

ANSWER: The effective bandwidth ranges from

$$500 \text{ sectors} \times 512 \text{ bytes/sector} \times (1/18.5 \text{ ms}) = 13.8 \text{ MB/s}$$

$$500 \text{ sectors} \times 512 \text{ bytes/sector} \times (1/16.4 \text{ ms}) = 15.6 \text{ MB/s}$$

This gives us a range of 13.8 MB/s / 54 MB/s = **26% to** 15.6 MB/s / 128 MB/s = **12%** of the maximum bandwidth from the inner to the outer tracks. □

So, even a fairly large request (500 sectors or 250 KB in this case) can incur significant overheads from seek and rotational latency.

EXAMPLE: Efficient access. For the disk described in Figure 12.3, how large must a request that begins on a random disk sector be to ensure that the disk gets at least 80% of its advertised maximum surface transfer bandwidth?

ANSWER: When reading a long sequence of logically sequential blocks, the disk will read an entire track, then do a 1 track seek (or a head switch and resettle, which amounts to the same thing) and then read the next track. Notice that track buffering allows the disk to read an entire track in one rotation regardless of which sector the head is over when it settles on the track and starts successfully reading. So, for the outer tracks, it reads for one rotation (8.4 ms) and then does a minimum seek (1 ms).

Thus, to achieve 80% of peak bandwidth after a random seek (10.5 ms), we need to read enough rotations worth of data to ensure that we spend 80% of the total time reading. If x is the number of rotations we will read, then we have:

$$0.8 \text{ totalTime} = x \text{ rotationTime}$$

$$0.8(10.5 \text{ ms} + (1 + 8.4)\, x \text{ ms}) = 8.4\, x \text{ ms}$$

$$x = 9.09$$

We therefore need to read at least 9.09 rotations worth of data to reach an efficiency of 80%. Since each rotation takes 8.4 ms and transfers data at 128 MB/s, 9.09 rotations transfers **9.77 MB** of data, **or** about **19,089 sectors**. □

12.1.3 Disk Scheduling

Because moving the disk arm and waiting for the platter to rotate is so expensive, performance can be significantly improved by optimizing the order in which pending requests are serviced. Disk scheduling can be done by the operating system, by the disk's firmware, or both.

FIFO. The simplest thing to do is to process requests in first-in-first-out (FIFO) order. Unfortunately, a FIFO scheduler can yield poor performance. For example, a sequence of requests that alternate between the outer and inner tracks of a disk will result in many long seeks.

SPTF/SSTF. An initially appealing option is to use a greedy scheduler that, given the current position of the disk head and platter, always services the pending request that can be handled in the minimum amount of time. This approach is called *shortest positioning time first* (SPTF) (or *shortest seek time first* (SSTF) if rotational positioning is not considered.)

SPTF and SSTF have two significant limitations. First, because moving the disk arm and waiting for some rotation time affects the cost of serving subsequent requests, these greedy approaches are not guaranteed to optimize disk performance. Second, these greedy approaches can cause starvation when, for example, a continuous stream of requests to inner tracks prevents requests to outer tracks from ever being serviced.

EXAMPLE: SPTF is not optimal. Suppose a disk's head is just inside the middle track of a disk so that seeking to the inside track would cost 9.9 ms while seeking to the outside track would cost 10.1 ms. Assume that for the disk in question, seeking between the outer and inner track costs 15 ms and that a rotation takes 10 ms.

Also suppose that the disk has two sets of pending requests. The first set is 1000 requests to read each of the 1000 sectors on the inner track of the disk; the second set is 2000 requests to read each of the 2000 sectors on the outer track of the disk.

Compare the average response time per request for the SPTF schedule (first read the "nearby" inner track and then read the outer track) and the alternative of reading the outer track first and then the inner track.

ANSWER: The total amount of time taken to complete all requests is slightly shorter by seeking first to the inside track and then to the outside. However, the *average* response time per request is less in the opposite order.

To service either set of requests, the disk must seek to the appropriate track and then wait for one full rotation while all of the track's data sweeps under the arm. For either set, the average response time for a request in that set will be the delay until the seek completes plus one half the disk's rotation time. Notice that the set handled second must wait until the first one is completely done before it can start, adding to the response time observed for requests in that set.

If we follow SPTF and read the sectors on the inner track first, the response time of the average request is the weighted average of the response time of the inner requests and the outer requests:

$$(1000\ (9.9 + 5) + 2000\ (9.9 + 10 + 15 + 5)) / 3000 = 31.6 \text{ ms}$$

If, instead, we read the outer tracks first, the weighted average is:

(2000 (10.1 + 5) + 1000 (10.1 + 10 + 15 + 5)) / 3000 = 23.3 ms

In this case, seeking to the nearest sector moves the disk head away from the majority of the requests, **increasing the overall average response time by 8.3 ms**. □

Elevator, SCAN, and CSCAN. Elevator-based algorithms like SCAN and CSCAN have good performance and also ensure fairness in that no request is forced to wait for an inordinately long time. The basic approach is similar to how an elevator works: when an elevator is going up, it keeps going up until all pending requests to go to floors above it have been satisfied; then, when an elevator is going down, it keeps going down until all pending requests to go to floors below it have been satisfied.

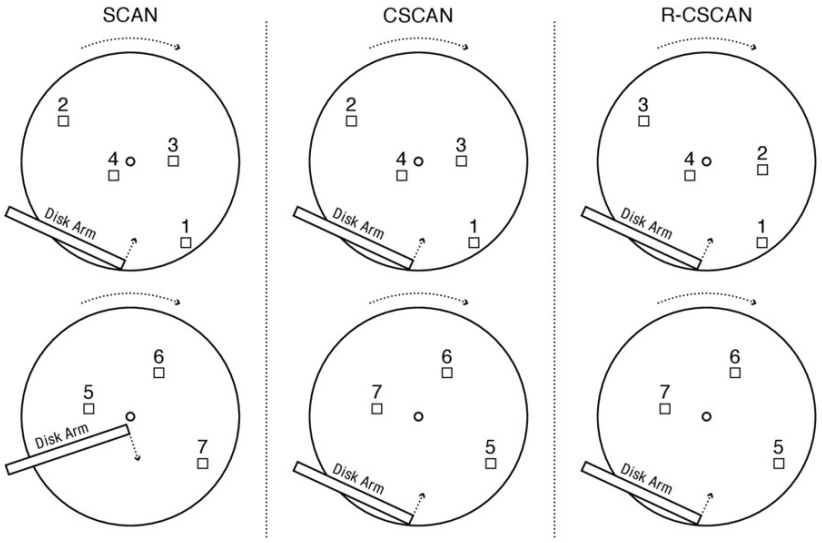


Figure 12.4: Elevator-based scheduling algorithms: (left) SCAN, (center) CSCAN, and (right) R-CSCAN. The numbering represents the order that each algorithm services requests 1-4 and then 5-7.

The *SCAN* scheduler works in the same way. The disk arm first sweeps from the inner to the outer tracks, servicing all requests that are between the arm's current position and the outer edge of the disk. Then, the arm sweeps from the outer to the inner tracks. Then the process is repeated. Figure 12.4-(left) illustrates the SCAN algorithm travelling from outer-to-inner tracks to service four pending requests and then travelling from inner-to-outer tracks to service three additional requests.

The *CSCAN* (circular SCAN) scheduler is a slight variation on SCAN in which the disk only services requests when the head is traveling in one direction (e.g., from inner tracks to outer ones). When the last request in the direction of travel is reached, the disk immediately seeks to where it started (e.g., the most inner track or the most inner track with a pending request) and services pending requests by moving the head *in the same direction* as the original pass (e.g., from inner tracks to outer ones again.) Figure 12.4-(center) illustrates the CSCAN algorithm travelling from outer-to-inner tracks to service four pending requests and then skipping to the outer track and travelling from outer-to-inner tracks to service three additional requests.

The advantage of CSCAN over SCAN is that if after a pass in one direction, the disk head were to just switch directions (as in SCAN), it will encounter a region of the disk where pending requests are sparse (since this region of the disk was just serviced). Seeking to the opposite side of the disk (as in CSCAN) moves the disk head to an area where pending requests are likely to be denser. In addition, CSCAN is more fair than SCAN in that seeking to the opposite side of the disk allows it to begin servicing the requests that likely been waiting longer than requests near but "just behind" the head.

Rather than pure seek-minimizing SCAN or CSCAN, schedulers also take into account rotation time and allow small seeks "in the wrong direction" to avoid extra rotational delays using the rotationally-aware *R-SCAN* or *R-CSCAN* variations. For example, if the disk head is currently over sector 0 of track 0 and there are pending requests at sector 1000 of track 0, sector 500 of track 1, and sector 0 of track 10,000, a R-CSCAN scheduler might service the second request, then the first, and then the third. Figure 12.4-(right) illustrates the R-CSCAN algorithm handling a request on the outer track, then one a few tracks in, then another request on the outer track, and a request near the center on the arm's first sweep. The arm's second sweep is the same as for CSCAN.

EXAMPLE: Effect of disk scheduling. For the disk described in Figure 12.3, consider a workload consisting of 500 read requests, each to a randomly chosen sector on disk, assuming that the disk head is on the outside track and that requests are serviced in CSCAN order from outside to inside. How long will servicing these requests take?

ANSWER: Answering a question like this requires making some educated guesses; different people may come up with different reasonable estimates here.

Seek time. We first note that with 500 pending requests spread randomly across the disk, the average seek from one request to the next will seek 0.2% of the way across the disk. With four surfaces, most of these seeks will also require a head switch. We don't know the exact time for a seek 0.2% of the way across the disk, but we can estimate it by interpolating between the time for a 1 track seek (1 ms) and the time for a 33.3% seek (10.5 ms for reads.) (Disk seek time is not actually linear in distance, but as we will see in a moment, the exact seek time seems unlikely to affect our answer much.)

estimated .2% seek time = (1 + 10.5 × .2/33.3) ms

= 1.06 ms

Rotation time. Since we don't know the position of the disk when the seek finishes and since sectors are scattered randomly, a simple and reasonable estimate for the time after the seek finishes for the desired block to rotate to the disk head is 4.15 ms, one half of the time that it takes a 7200 RPM disk to rotate once.

Transfer time. Similar to the example with FIFO servicing of the same requests, the transfer time for each sector is at most 0.0095 ms.

Total time. 1.06 + 4.15 + .0095 = 5.22 ms per request, so **500 requests will take about 2.6 s**. Notice that the time for the SCAN scheduled time is less than half the 7.8 s time for FIFO servicing of the same requests. □

12.2 Flash Storage

Over the past decade, flash storage has become a widely used storage medium. Flash storage is the dominant storage technology for handheld devices from phones to cameras to thumb drives, and it is used in an increasing fraction of laptop computers and machine room servers.

Flash storage is a type of *solid state storage*: it has no moving parts and stores data using electrical circuits. Because it has no moving parts, flash storage can have much better random IO performance than disks, and it can use less power and be less vulnerable to physical damage. On the other hand, flash storage remains significantly more expensive per byte of storage than disks.

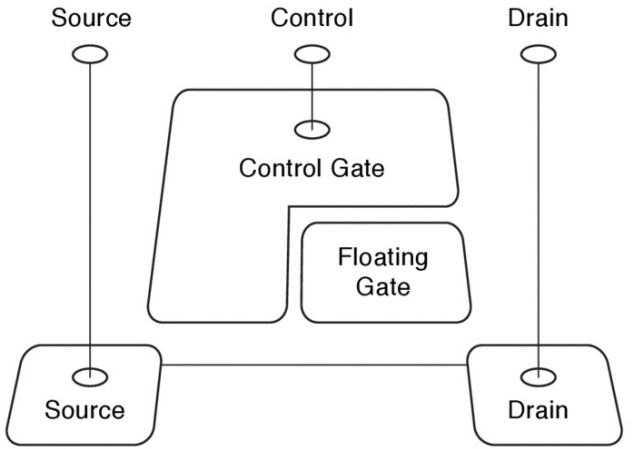


Figure 12.5: A floating gate transistor.

Each flash storage element is a floating gate transistor. As Figure 12.5 illustrates, an extra gate in such a transistor "floats" — it is not connected to any circuit. Since the floating gate is entirely surrounded by an insulator, it will hold an electrical charge for months or years without requiring any power. Even though the floating gate is not electrically connected to anything, it can be charged or discharged via electron tunneling by running a sufficiently high-voltage current near it. The floating gate's state of charge affects the transistor's threshold voltage for activation. Thus, the floating gate's state can be detected by applying an intermediate voltage to the transistor's control gate that will only be sufficient to activate the transistor if the floating gate is changed.

In single-level flash storage, the floating gate stores one bit (charge or not charged); in multi-level flash storage, the floating gate stores multiple bits by storing one of several different charge levels.

NOR flash storage is wired to allow individual words to be written and read. NOR flash storage is useful for storing device firmware since it can be executed in place. NAND flash storage is wired to allow reads and writes of a page at a time, where a page is typically 2 KB to 4 KB. NAND flash is denser than NOR flash, so NAND is used in the storage systems we will consider.

Flash storage access and performance. Flash storage is accessed using three operations.

- **Erase erasure block.** Before flash memory can be written, it must be erased by setting each cell to a logical "1". Flash memory can only be erased in large units called *erasure blocks*. Today, erasure blocks are often 128 KB to 512 KB. Erasure is a slow operation, usually taking several milliseconds.

  Erasing an erasure block is what gives flash memory its name for its resemblance to the flash of a camera.

- **Write page.** Once erased, NAND flash memory can be written on a page-by-page basis, where each page is typically 2048-4096 bytes. Writing a page typically takes tens of microseconds.
- **Read page.** NAND flash memory can be read on a page by page basis. Reading a page typically takes tens of microseconds.

Notice that to write a page, its entire erasure block must first be erased. This is a challenge both because erasure is slow and because erasure affects a large number of pages. Flash drives implement a *flash translation layer* (FTL) that maps logical flash pages to different physical pages on the flash device. Then, when a single logical page is overwritten, the FTL writes the new version to some free, already-erased physical page and remaps the logical page to that physical page.

Write remapping significantly improves flash performance.

EXAMPLE: Remapping flash writes. Consider a flash drive with a 4 KB pages, 512 KB erasure blocks, 3 ms flash times, and 50 µs read-page and write-page times. Suppose writing a page is done with a naive algorithm that reads an entire erasure block, erases it, and writes the modified erasure block. How long would each page write take?

ANSWER: This naive approach would require:

((512 KB/erasure block) / 4 KB/page) × (page read time + page write time) + block erase time

= 128 × (50 + 50) µs + 3 ms

= **15.8 ms per write**

□

Suppose remapping is used and that a flash device always has at least one unused erasure block available for a target workload. How long does an average write take now?

ANSWER: With remapping, the cost of flashing an erasure block is amortized over 512/4 = 128 page writes. This scenario gives a cost of (3 ms⁄128) + 50 µs = **73.4 µs** per write. □

In practice, there is likely to be some additional cost per write under the remapping scheme because in order to flash an erasure block to free it for new writes, the firmware may need to garbage collect live pages from that erasure block and copy those live pages to a different erasure block.

Internally, a flash device may have multiple independent data paths that can be accessed in parallel. Therefore, to maximize sustained bandwidth when accessing a flash device, operating systems issue multiple concurrent requests to the device.

Durability. Normally, flash memory can retain its state for months or years without power. However, over time the high current loads from flashing and writing memory causes the circuits to degrade. Eventually, after a few thousand to a few million program-erase cycles (depending on the type of flash), a given cell may *wear out* and no longer reliably store a bit.

In addition, reading a flash memory cell a large number of times can cause the surrounding cells' charges to be disturbed. A *read disturb error* can occur if a location in flash memory is read to many times without the surrounding memory being written.

To improve durability in the face of wear from writes and disturbs from reads, flash devices make use of a number of techniques:

- **Error correcting codes.** Each page has some extra bytes that are used for error correcting codes to protect against bit errors in the page.
- **Bad page and bad erasure block management.** If a page or erasure block has a manufacturing defect or wears out, firmware on the device marks it as bad and stops storing data on it.
- **Wear leveling.** As noted above, rather than overwrite a page in place, the flash translation layer remaps the logical page to a new physical page that has already been erased. This remapping ensures that a hot page that is overwritten repeatedly does not prematurely wear out a particular physical page on the flash device.

  Wear leveling moves a flash device's logical pages to different physical pages to ensure that no physical page gets an inordinate number of writes and wears out prematurely. Some wear leveling algorithms also migrate unmodified pages to protect against read disturb errors.
- **Spare pages and erasure blocks.** Flash devices can be manufactured with spare pages and spare erasure blocks in the device. This spare capacity serves two purposes.

  First, it provides extra space for wear leveling: even if the device is logically "full" the wear leveling firmware can copy live pages out of some existing erasure blocks into a spare erasure block, allowing it to flash those existing erasure blocks.

  Second, it allows bad page and bad erasure block management to function without causing the logical size of the device to shrink.

In addition to affecting reliability, wear out affects a flash device's performance over time.

First, as a device wears out, accesses may require additional retries, slowing them. Second, as spare pages and erasure blocks are consumed by bad ones, the wear leveling algorithms have less spare space and have to garbage collect live pages — copying them out of their existing erasure blocks — more frequently.

	Size
Capacity	300 GB

Page Size	4 KB
Performance	
Bandwidth (Sequential Reads)	270 MB/s
Bandwidth (Sequential Writes)	210 MB/s
Read/Write Latency	75 µs
Random Reads Per Second	38,500
Random Writes Per Second	2,000
	2,400 with 20% space reserve
Interface	SATA 3 Gb/s
Endurance	
Endurance	1.1 PB
	1.5 PB with 20% space reserve
Power	
Power Consumption Active/Idle	3.7 W / 0.7 W

Figure 12.6: Key parameters for an Intel 710 Series Solid State Drive manufactured in 2011.

Example: Intel 710 Series Solid-State Drive. Figure 12.6 shows some key parameters for an Intel 710 Series solid state drive manufactured in 2011. This drive uses multi-level NAND flash to get high storage densities. Normally, multi-level flash is less durable than single-level, but this Intel drive uses sophisticated wear leveling algorithms and a large amount of spare space to provide high durability.

The sequential performance of this drive is very good, with peak sustained read and write bandwidths of 270 MB/s and 210 MB/s respectively. In comparison, a high-end Seagate Cheetah 15K.7 drive manufactured in 2010 spins at 15,000 revolutions per minute and provides 122 MB/s to 204 MB/s of sustained bandwidth.

Random read performance is excellent. The latency for a single random 4 KB read is just 75 µs, and when multiple concurrent requests are in flight, the drive can process 38,500 random reads per second — one every 26 µs. This is orders of magnitude better than the random read performance of a spinning disk drive.

Random write performance is also very good, but not as good as random read performance. The latency for a single random 4 KB write is 75 µs; the drive reduces write latency by buffering writes in volatile memory, and it has capacitors that store enough charge to write all buffered updates to flash storage if a power loss occurs.

When multiple concurrent writes are in flight, the drive can process 2,000 random writes per second when it is full; if it is less than 80% full, that number rises to 2,400. Random write throughput increases when the drive has more free space because the drive has to garbage collect live pages from erasure blocks less often and because when the drive eventually does do that garbage collection, the erasure blocks are less full.

The drive's is rated for 1.1 PB (1.1×10^{15} bytes) of endurance (1.5 PB if it is less than 80% full.) For many workloads, this endurance suffices for years or decades of use. However, solid state drives may not always be a good match for high-bandwidth write streaming. In the extreme, an application constantly streaming writes at 200 MB/s could wear this drive out in 64 days.

Technology affects interfaces — the TRIM command

Historically, when a file system deleted a file stored on a spinning disk, all it needed to do was to update the file's metadata and the file system's free space bitmap. It did not need to erase or overwrite the file's data blocks on disk — once the metadata was updated, these blocks could never be referenced, so there was no need to do anything with them.

When such file systems were used with flash drives, users observed that their drives got slower over time. As the amount of free space fell, the drives' flash translation layer was forced to garbage collect erasure blocks more frequently; additionally, each garbage collection pass became more expensive because there were more live pages to copy from old erasure blocks to the new ones.

Notice that this slowing could occur even if the file system appeared to have a large amount of free space. For example, if a file system moves a large file from one range of blocks to another, the storage hardware has no way to know that the pages in the old range are no longer needed unless the file system can tell it so.

The TRIM command was introduced into many popular operating systems between 2009 and 2011 to allow file systems to inform the underlying storage when the file system has stopped using a page of storage. The TRIM command makes the free space known to the file system visible to the underlying storage layer, which can significantly reduce garbage collection overheads and help flash drives retain good performance as they age.

EXAMPLE: Random read workload. For the solid state disk described in Figure 12.6, consider a workload consisting of 500 read requests, each of a randomly chosen page. How long will servicing these requests take?

ANSWER: The disk can service random read requests at a rate of 38,500 per second, so

500 requests will take 500/38500 = **13 ms**. In contrast, for the spinning disk example, the same 500 requests would take 7.33 seconds. □

EXAMPLE: Random vs. sequential reads. How does this drive's random read performance compare to its sequential read performance?

ANSWER: The effective bandwidth in this case is 500 requests × (4 KB/request) / 13 milliseconds = 158 MB/s. The random read bandwidth is thus 158/270 = **59% of the sequential read bandwidth**. □

EXAMPLE: Random write workload. For the solid state disk described in Figure 12.6, consider a workload consisting of 500 write requests, each of a randomly chosen page. How long will servicing these requests take?

ANSWER: The disk can service random write requests at a rate of 2000 per second (assuming the disk is nearly full), so 500 requests will take 500/2000 = **250 ms**. □

EXAMPLE: Random vs. sequential writes. How does this random write performance compare to the drive's sequential write performance?

ANSWER: The effective bandwidth in this case is 500 requests × (4 KB/request) / 250 ms = 8.2 MB/s. The random write bandwidth is thus 8.2/210 = **3.9% of the sequential write bandwidth**. □

12.3 Summary and Future Directions

Today, spinning disk and flash memory dominate storage technologies, and each has sufficient advantages to beat the other for some workloads and environments.

Spinning disk vs. flash storage. Spinning disks are often used when capacity is the primary goal. For example, spinning disk is often used for storing media files and home directories. For workloads limited by storage capacity, spinning disks can often provide much better capacity per dollar than flash storage. For example, in October 2011, a 2 TB Seagate Barracuda disk targeted at workstations cost about $80 and a 300 GB Intel 320 Series solid state drive targeted at laptops cost about $600, giving the spinning disk about a 50:1 advantage in GB per dollar.

Both spinning disks and flash storage are viable when sequential bandwidth is the goal. In October 2011, flash drives typically have modestly higher per-drive sequential bandwidths than spinning drives, but the spinning drives typically have better sequential bandwidth per dollar spent than flash drives. For example, the same Seagate disk has a sustained bandwidth of 120 MB/s (1.5 MB/s per dollar) while the same Intel SSD has a read/write bandwidth of 270/205 MB/s (about 0.4 MB/s per dollar.)

Flash storage is often used when good random access performance or low power consumption is the goal. For example, flash storage is frequently used in database transaction processing servers, in smart phones, and in laptops. For example, the Seagate drive described above rotates at 5900 RPM, so it takes about 5 ms for a half rotation. Even with good scheduling and even if data is confined to a subset of tracks, it would be hard to get more than 200 random I/Os per second from this drive (about 2.5 random I/Os per second per dollar.) Conversely, the Intel SSD can sustain 23,000 random writes and 39,500 random reads per second (about 38 or 66 random writes or reads per second per dollar.)

With respect to power, spinning disks typically consume 10-20W depending on whether it is just spinning or actively reading and writing data, while a flash drive might consume 0.5W-1W when idle and 3-5W when being accessed. Flash drives' power advantage makes them attractive for portable applications such as laptop and smartphone storage.

Figure 12.7: In 2011, flash storage "keys" such as this one can store as much as 256 GB in a device that is a few centimeters long, and 1-2 cm wide and tall.

Flash memory can also have a significant form factor advantage with respect to physical size and weight. Although some flash drives are designed as drop-in replacements for spinning disks and so are similar in size, flash storage can be much smaller than a typical spinning disk. For example, in 2011, a USB flash storage "key" such as the one in Figure 12.7 can store as much as 256 GB in a device that is not much larger than a house key.

Metric	**Spinning Disk**	**Flash**
Capacity / Cost	Excellent	Good

Sequential BW / Cost	Good	Good
Random I/O per Second / Cost	Poor	Good
Power Consumption	Fair	Good
Physical Size	Good	Excellent

Figure 12.8: Relative advantages and disadvantages of spinning disk and flash storage.

Figure 12.8 summarizes these advantages and disadvantages; of course, many systems need to do well on multiple metrics, so system designers may need to compromise on some metrics or use combinations of technologies.

Technology trends. Over the past decades, the cost of storage capacity has fallen rapidly for both spinning disks and solid state storage. Compare the 2 TB disk drive for $80 in 2011 to a 15 MB drive costing $113 in 1984 (or about $246 in 2011 dollars): the cost per byte has improved by a factor of about 400,000 over 27 years — over 50% per year for nearly 3 decades.

The first disk drive

Prior to the invention of magnetic disks, magnetic cylinders, called drums, were used for on-line storage. These drums spun on their axes and typically had one head per track. So, there was no seek time to access a block of data; one merely waited for a block to rotate underneath its head.

By using spinning disks instead of drums, the magnetic surface area, and hence the storage capacity, could be increased.

The first disk drive, the IBM 350 Disk System (two are shown in the foreground of this photograph), was introduced in 1956 as part of the IBM RAMAC ("Random Access Method of Accounting and Control") 305 computer system. The 350 Disk system stored about 3.3 MB on 50 platters, rotated its platters at 1200 RPM, had an average seek time of 600 ms, and weighed about a ton. The RAMAC 305 computer system with its 350 disk system could be leased for $3,200 per month. Assuming a useful life of 5 years and converting to 2011 dollars, the cost was approximately $1.3 million for the system — about $400,000 per megabyte.

Recent rates of improvement for flash storage have been even faster. For example, in 2001, the Adtron S35PC 14 GB flash drive cost $42,000. Today's Intel 320 costs 70 times less for 21 times more capacity, an improvement of about 2x per year over the past decade.

Similar capacity improvements for spinning disk and flash are expected for at least the next few years. Beyond that, there is concern that we will be approaching the physical limits of both magnetic disk and flash storage, so the longer-term future is less certain. (That said, people have worried that disks were approaching their limits several times in the past, and we will not be surprised if the magnetic disk and flash industries continue rapid improvements for quite a few more years.)

In contrast to capacity, performance is likely to improve more slowly for both technologies. For example, a mid-range spinning disk in 1991 might have had a maximum bandwidth of 1.3 MB/s and an average seek time of 17 ms. Bandwidths have improved by about a factor of 90 in two decades (about 25% per year) while seek times and rotational latencies have only improved by about a factor of two (less than 4% per year.) Bandwidths have improved more quickly than rotational latency and seek times because bandwidth benefits from increasing storage densities, not just increasing rotational rates.

For SSDs, the story is similar, though recent increases in volumes have helped speed the pace of improvements. For example, in 2006 a BitMicro E-Disk flash drive could provide 9,500 to 11,700 random reads per second and 34-44 MB/s sustained bandwidth. Compared to the Intel 320 SSD from 2011, bandwidths have improved by about 40% per year and random access throughput has improved by about 25% per year over the past 5 years.

New technologies. This is an exciting time for persistent storage. After decades of undisputed reign as the dominant technology for on-line persistent storage, spinning magnetic disks are being displaced by flash storage in many application domains, giving both operating system designers and application writers an opportunity to reexamine how to best use storage. Looking forward, many researchers speculate that new technologies may soon be nipping at the heels and even surpassing flash storage.

For example, *phase change memory (PCM)* uses a current to alter the state of chalcogenide glass between amorphous and crystalline forms, which have significantly different electrical resistance and can therefore be used to represent data bits. Although PCM does not yet match the density of flash, researchers speculate that the technology is fundamentally more scalable and will ultimately be able to provide higher storage densities at lower costs. Furthermore, PCM is expected to have much better write performance and endurance than flash.

As another example, a *memristor* is a circuit element whose resistance depends on the amounts and directions of currents that have flowed through it in the past. A number of different memristor constructions are being pursued, and some have quite promising properties. For example, in 2010 Hewlett Packard labs described a prototype memristor constructed of a thin titanium dioxide film with 3 nm by 3 nm storage elements that can switch states in 1 ns. These densities are similar to contemporary flash memory devices and these switching times are similar to contemporary DRAM chips. The devices also

have write endurance similar to flash, and extremely long (theoretically unlimited) storage lifetimes. Furthermore, researchers believe that these and others memristors' densities will scale well in the future. For example, a design for 3-D stacking of memristors was published in 2009 in the *Proceedings of National Academy of Sciences* by Dmitri Strukov and R. Stanley Williams of HP Labs. http://www.pnas.org/content/106/48/20155.abstract

If technologies such as these pan out as hoped, operating system designers will have opportunities to rethink our abstractions for both volatile and nonvolatile storage: how should we make use of word-addressable, persistent memory with densities exceeding current flash storage devices and with memory access times approaching those of DRAM? What could we do if each core on a 32 core processor chip had access to a few gigabytes of stacked memristor memory?

Exercises

Size	
Form factor	2.5 inch
Capacity	320 GB
Performance	
Spindle speed	5400 RPM
Average seek time	12.0 ms
Maximum seek time	21 ms
Track-to-track seek time	2 ms
Transfer rate (surface to buffer)	850 Mbit/s (maximum)
Transfer rate (buffer to host)	3 Gbit/s
Buffer memory	8 MB

Figure 12.9: Hardware specifications for a 320 GB SATA disk drive.

1. **Discussion.** Some high-end disks in the 1980s had multiple disk arm assemblies per disk enclosure in order to allow them to achieve higher performance. Today, high-performance server disks have a single arm assembly per disk enclosure. Why do you think disks so seldom have multiple disk arm assemblies today?
2. For the disk described in Figure 12.3:
   a. What is the range of the number of sectors per track on the disk?
   b. Estimate the number of tracks on the disk.
   c. Estimate the distance from the center of one track to the center of the next track.
3. A disk may have multiple surfaces, arms, and heads, but when you issue a read or write, only one head is active at a time. It seems like one could greatly increase disk bandwidth for large requests by reading or writing with all of the heads at the same time. Given the physical characteristics of disks, can you figure out why no one does this?
4. For the disk described in Figure 12.3, consider a workload consisting of 500 read requests, each of a randomly chosen sector on disk, assuming that the disk head is on the outside track and that requests are serviced in P-CSCAN order from outside to inside. How long will servicing these requests take?

   Note: Answering this question will require making some estimates.
5. Suppose I have a disk such as the 320 GB SATA drive described in Figure 12.9 and I have a workload consisting of 10000 reads to sectors randomly scattered across the disk. How long will these 10000 requests take (total) assuming the disk services requests in FIFO order?
6. Suppose I have a disk such as the 320 GB SATA drive described in Figure 12.9 and I have a workload consisting of 10000 reads to 10000 sequential sectors on the outer-most tracks of the disk. How long will these 10000 requests take (total) assuming the disk services requests in FIFO order?
7. Suppose I have a disk such as the 320 GB SATA drive described in Figure 12.9 and I have a workload consisting of 10000 reads to sectors randomly scattered across the disk. How long will these 10000 requests take (total) assuming the disk services requests using the SCAN/Elevator algorithm.
8. Suppose I have a disk such as the 320 GB SATA drive described in Figure 12.9 and I have a workload consisting of 10000 reads to sectors randomly scattered across a 100 MB file, where the 100 MB file is laid out sequentially on the disk. How long will these 10000 requests take (total) assuming the disk services requests using the SCAN/Elevator algorithm?
9. Write a program that creates a 100 MB file on your local disk and then measures the time to do each of four things:
   a. **Sequential overwrite.** Overwrite the file with 100 MB of new data by writing the file from beginning to end and then calling fsync() (or the equivalent on your platform).
   b. **Random buffered overwrite.** Do the following 50,000 times: choose a 2 KB-aligned offset in the file uniformly at random, seek to that location in the file, and write 2 KB of data at that position. Then, once all 50,000 writes have been issued, call fsync() (or the equivalent on your platform).

c. **Random buffered overwrite.** Do the following 50,000 times: choose a 2 KB-aligned offset in the file uniformly at random, seek to that location in the file, write 2 KB of data at that position, and call fsync() (or the equivalent on your platform) after each individual write.

d. **Random read.** Do the following 50,000 times: choose a 2 KB-aligned offset in the file uniformly at random, seek to that location in the file, and read 2 KB of data at that position.

Explain your results.

10. Write a program that creates three files, each of 100 MB, and then measures the time to do each of three things:

    a. **fopen()/fwrite().** Open the first file using fopen() and issue 256,000 sequential four-byte writes using fwrite().

    b. **open()/write().** Open the second file using open() and issue 256,000 sequential four-byte writes using write().

    c. **mmap()/store.** Map the third file into your program's memory using mmap() and issue 256,000 sequential four-byte writes by iterating through memory and writing to each successive word of the mapped file.

    Explain your results.

Size	
Usable capacity	2 TB (SLC flash)
Cache Size	64 GB (Battery-backed RAM)
Page Size	4 KB
Performance	
Bandwidth (Sequential Reads from flash)	2048 MB/s
Bandwidth (Sequential Writes to flash)	2048 MB/s
Read Latency (cache hit)	15 µs
Read Latency (cache miss)	200 µs
Write Latency	15 µs
Random Reads (sustained from flash)	100,000 per second
Random Writes (sustained to flash)	100,000 per second
Interface	8 Fibre Channel ports with 4 Gbit/s per port
Power	
Power Consumption	300 W

Figure 12.10: Key parameters for a hypothetical high-end flash drive in 2011.

---

11. Suppose that you have a 256 GB solid state drive that the operating system and drive both support the TRIM command. To evaluate the drive, you do an experiment where you time the system's write performance for random page-sized when the file system is empty compared to its performance when the file system holds 255 GB of data, and you find that write performance is significantly worse in the latter case.

    What is the likely reason for this worse performance as the disk fills despite its support for TRIM?

    What can be done to mitigate this slowdown?

12. Suppose you have a flash drive such as the one described in Figure 12.10 and you have a workload consisting of 10000 4 KB reads to pages randomly scattered across the drive. Assuming that you wait for request i to finish before you issue request i + 1, how long will these 10000 requests take (total)?

13. Suppose you have a flash drive such as the one described in Figure 12.10 and you have a workload consisting of 10000 4 KB reads to pages randomly scattered across the drive. Assuming that you issue requests concurrently, using many threads, how long will these 10000 requests take (total)?

14. Suppose you have a flash drive such as the one described in Figure 12.10 and you have a workload consisting of 10000 4 KB writes to pages randomly scattered across the drive. Assuming that you wait for request i to finish before you issue request i + 1, how long will these 10000 requests take (total)?

15. Suppose you have a flash drive such as the one described in Figure 12.10 and you have a workload consisting of 10000 4 KB writes to pages randomly scattered across the drive. Assuming there are a large number of threads to issue writes concurrently, how long will these 10000 requests take (total)?

16. Suppose you have a flash drive such as the one described in Figure 12.10 and you have a workload consisting of 10000 4 KB reads to 10000 sequential pages. How long will these 10000 request take (total)?

13. Files and Directories

What's in a name? That which we call a rose
By any other name would smell as sweet. —*Juliet*
Romeo and Juliet (II, ii, 1-2)
(Shakespeare)

Recall from Chapter 11 that file systems use directories to provide hierarchically named files, and that each file contains metadata and a sequence of data bytes. However, as Chapter 12 discussed, storage devices provide a much lower-level abstraction — large arrays of data blocks. Thus, to implement a file system, we must solve a translation problem: How do we go from a file name and offset to a block number?

A simple answer is that file systems implement a dictionary that maps keys (file name and offset) to values (block number on a device). We already have many data structures for implementing dictionaries, including hash tables, trees, and skip lists, so perhaps we can just use one of them?

Unfortunately, the answer is not so simple. File system designers face four major challenges:

- **Performance.** File systems need to provide good performance while coping with the limitations of the underlying storage devices. In practice, this means that file systems strive to ensure good *spatial locality*, where blocks that are accessed together are stored near one another, ideally in sequential storage blocks.
- **Flexibility.** One major purpose of file systems is allowing applications to share data, so file systems must be jacks-of-all-trades. They would be less useful if we had to use one file system for large sequentially-read files, another for small seldom-written files, another for large random-access files, another for short-lived files, and so on.
- **Persistence.** File systems must maintain and update both user data and their internal data structures on persistent storage devices so that everything survives operating system crashes and power failures.
- **Reliability.** File systems must be able to store important data for long periods of time, even if machines crash during updates or some of the system's storage hardware malfunctions.

This chapter discusses how file systems are organized to meet the first three challenges. Chapter 14 addresses reliability.

13.1 Implementation Overview

File systems must map file names and offsets to physical storage blocks in a way that allows efficient access. Although there are many different file systems, most implementations are based on four key ideas: directories, index structures, free space

maps, and locality heuristics.

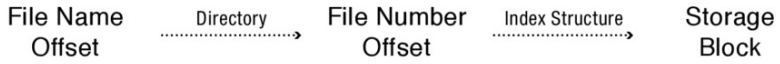


Figure 13.1: File systems map file names and file offsets to storage blocks in two steps. First, they use directories to map names to file numbers. Then they use an index structure such as a persistently stored tree to find the block that holds the data at any specific offset in that file.

Directories and index structures. As Figure 13.1 illustrates, file systems map file names and file offsets to specific storage blocks in two steps.

First, they use *directories* to map human-readable file names to file numbers. Directories are often just special files that contain lists of *file name* → *file number* mappings.

Second, once a file name has been translated to a file number, file systems use a persistently stored *index structure* to locate the blocks of the file. The index structure can be any persistent data structure that maps a file number and offset to a storage block. Often, to efficiently support a wide range of file sizes and access patterns, the index structure is some form of tree.

Free space maps. File systems implement *free space maps* to track which storage blocks are free and which are in use as files grow and shrink. At a minimum, a file system's free space map must allow the file system to find a free block when a file needs to grow, but because spatial locality is important, most modern file systems implement free space maps that allow them to find free blocks near a desired location. For example, many file systems implement free space maps as bitmaps in persistent storage.

Locality heuristics. Directories and index structures allow file systems to locate desired file data and metadata no matter where they are stored, and free space maps allow them to locate the free space near any location on the persistent storage device. These *mechanisms* allow file systems to employ various *policies* to decide where a given block of a given file should be stored.

These policies are embodied in *locality heuristics* for grouping data to optimize performance. For example, some file systems group each directory's files together but spread different directories to different parts of the storage device. Others periodically *defragment* their storage, rewriting existing files so that each file is stored in sequential storage blocks and so that the storage device has long runs of sequential free space so that new files can be written sequentially. Still others optimize writes over reads and write all data sequentially, whether a given set of writes contains updates to one file or to many different ones.

Implementation details. In this chapter, we first discuss how directories are implemented. Then, we look at the details of how specific file systems handle the details of placing and finding data in persistent storage by implementing different index structures, free space maps, and locality heuristics.

13.2 Directories: Naming Data

As Figure 13.1 indicates, to access a file, the file system first translates the file's name to its number. For example, the file called /home/tom/foo.txt might internally known as file 66212871. File systems use directories to store their mappings from human-readable names to internal file numbers, and they organize these directories hierarchically so that users can group related files and directories.

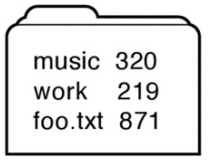


Figure 13.2: A directory is a file that contains a collection of *file name* → *file number* mappings.

Implementing directories in a way that provides hierarchical, name-to-number mappings turns out to be simple: use files to store directories. So, if the system needs to determine a file's number, it can just open up the appropriate directory file and scan through the file name/file number pairs until it finds the right one.

For example, illustrates Figure 13.2 the contents of a single directory file. To open file foo.txt, the file system would scan this directory file, find the foo.txt entry, and see that file foo.txt has file number 66212871.

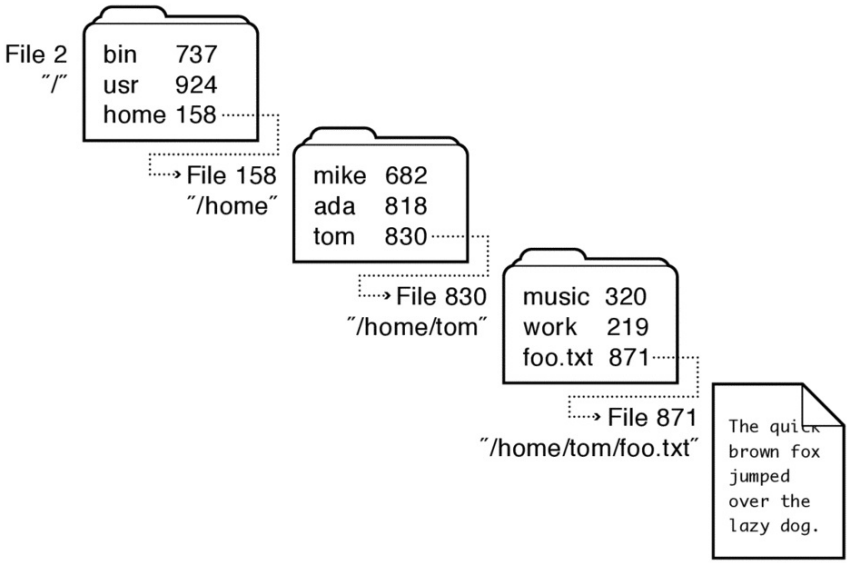


Figure 13.3: Directories can be arranged hierarchically by having one directory contain the *file name → file number* mapping for another directory.

Of course, if we use files to store the contents of directories such as /home/tom, we still have the problem of finding the directory files, themselves. As Figure 13.3 illustrates, the file number for directory /home/tom can be found by looking up the name tom in the directory /home, and the file number for directory /home can be found by looking up the name home in the root directory /.

Recursive algorithms need a base case — we cannot discover the root directories file number by looking in some other directory. The solution is to agree on the root directory's file number ahead of time. For example, the Unix Fast File System (FFS) and many other Unix and Linux file systems use two as the predefined file number for the root directory of a file system.

So, to read file /home/tom/foo.txt in Figure 13.3, we first read the root directory by reading the file with the well-known root number two. In that file, we search for the name home and find that directory /home is stored in file 88026158. By reading file 88026158 and searching for the name tom, we learn that directory /home/tom is stored in file 5268830. Finally, by reading file 5268830 and searching for the name foo.txt, we learn that /home/tom/foo.txt is file number 66212871.

Although looking up a file's number can take several steps, we expect there to be locality (e.g., when one file in a directory is accessed, other files in the directory are often likely to be accessed soon), so we expect that caching will reduce the number of disk accesses needed for most lookups.

Directory API. If file systems use files to store directory information, can we just use the standard open/close/read/write API to access them?

No. Directories use a specialized API because they must control the contents of these files. For example, file systems must prevent applications from corrupting the list of *name → file number* mappings, which could prevent the operating system from performing lookups or updates. As another example, the file system should enforce the invariant that each file number in a valid directory entry refers to a file that actually exists.

File systems therefore provide special system calls for modifying directory files. For example, rather than using the standard write system call to add a new file's entry to a directory, applications use the create call. By restricting updates, these calls ensure that directory files can always be parsed by the operating system. These calls also bind together the creation or removal of a file and the file's directory entry, so that directory entries always refer to actual files and that all files have at least one directory entry.

In the API described in Chapter 11, the other calls that modify directory files are mkdir, link, unlink, and rmdir.

So, for example, for the file system illustrated in Figure 13.3, Tom could rename foo.txt to hw1.txt in his home directory by running a process that makes the following two system calls

```
link(''foo.txt'', ''hw1.txt'');
unlink(''foo.txt'');
```

Processes can simply read directory files with the standard read call.

EXAMPLE: Reading directories. It is useful for programs to be able to get a list of all file names in a directory to, for example, recursively traverse a hierarchy from some point. However, the file system API described in Chapter 11 does not have call specifically for reading directories.

Given just the system call API in that figure, how could a process learn the names of files in the process's current working directory?

ANSWER: Processes can read the contents of directory files using the standard file read system call used to read the contents of "normal" files.

Although operating systems must restrict writes to directory files to ensure invariants on directory structure, they need not restrict applications from reading the contents of directory files (that they have permission to read). For simplicity, applications would access this function via a standard library that also includes routines for parsing directory files. □

Although it is not fundamentally necessary to have dedicated system calls for reading directories, it can be convenient. For example, Linux includes a getdents ("get directory entries") system call that reads a specified number of directory entries from an open file.

Directory internals. Many early implementations simply stored linear lists of *file name*, *file number* pairs in directory files. For example, in the original version of the Linux ext2

file system, each directory file stored a linked list of directory entries as illustrated in Figure 13.4.

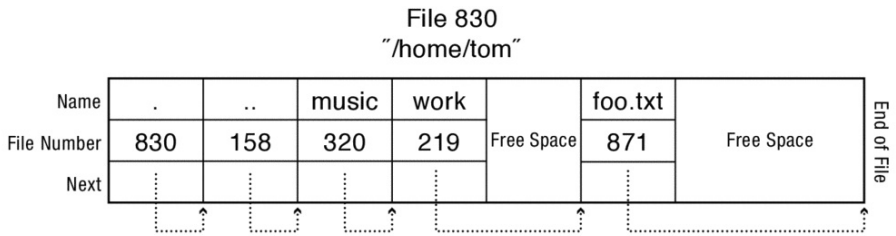
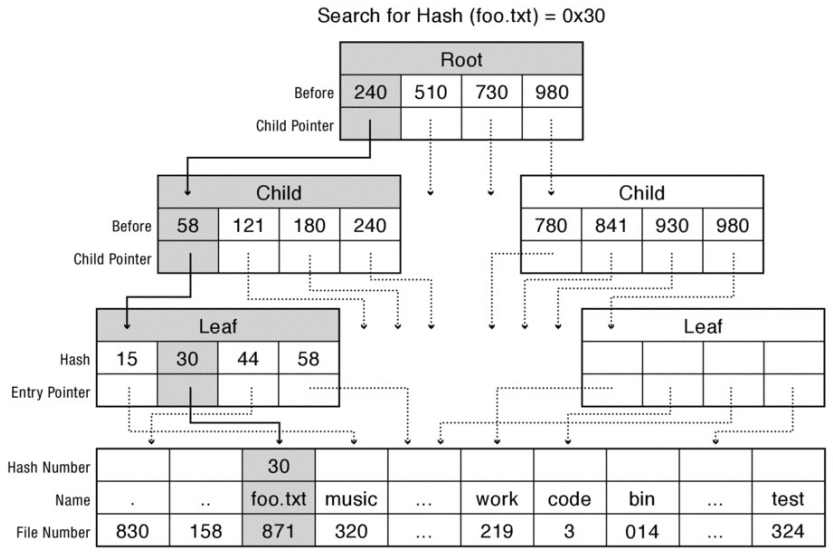


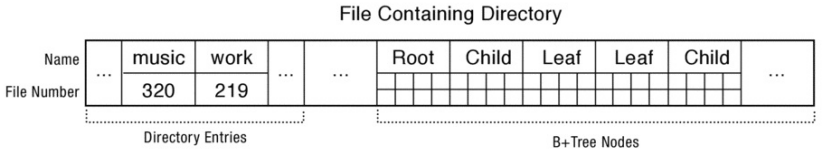
Figure 13.4: A linked list implementation of a directory. This example shows a directory file containing five entries: Music, Work, and foo.txt, along with . (the current directory) and .. (the parent directory).

Simple lists work fine when the number of directory entries is small, and that was the expected case for many early file systems, but systems occasionally encounter workloads that generate thousands of files in a directory. Once a directory has a few thousand entries, simple list-based directories become sluggish.

To efficiently support directories with many entries, many recent file systems including Linux XFS, Microsoft NTFS, and Oracle ZFS organize a directory's contents as a tree. Similarly, newer versions of ext2 augment the underlying linked list with an additional hash-based structure to speed searches.



(a) Logical view



(b) Physical storage

Figure 13.5: Tree-based directory structure similar to the one used in Linux's XFS file system.

For example, Figure 13.5-(a) illustrates a tree-based directory structure similar to the one used in Linux XFS, and Figure 13.5-(b) illustrates how these records are physically arranged in a directory file.

In this example, directory records mapping file names to file numbers are stored in a B+tree that is indexed by the hash of each file's name. To find the file number for a given file name (e.g., out2), the file system first computes a hash of the name (e.g., 0x0000c194). It then uses that hash as a key to search for the directory entry in the tree:

starting at the B+tree root at a well-known offset in the file (BTREE_ROOT_PTR), and proceeding through the B+tree's internal nodes to the B+tree's leaf nodes. At each level, a tree node contains an array of *(hash key, file offset)* pairs that each represent a pointer to the child node containing entries with keys smaller than *hash key* but larger than the previous entry's *hash key*. The file system searches the node for the first entry with a *hash key* value that exceeds the target key, and then it follows the corresponding *file offset* pointer to the correct child node. The *file offset* pointer in the record at the leaf nodes points to the target directory entry.

In the XFS implementation, directory entries are stored in the first part of the directory file. The B+tree's root is at a well-known offset within file (e.g., BTREE_ROOT_PTR). The fixed-size internal and leaf nodes are stored after the root node, and the variable-size directory entries are stored at the start of the file. Starting from the root, each tree node includes pointers to where in the file its children are stored.

Hard and soft links. Many file systems allow a given file to have multiple names. For example, /home/tom/Work/Classes/OS/hw1.txt and /home/tom/todo/hw1.txt may refer to the same file, as Figure 13.6 illustrates.

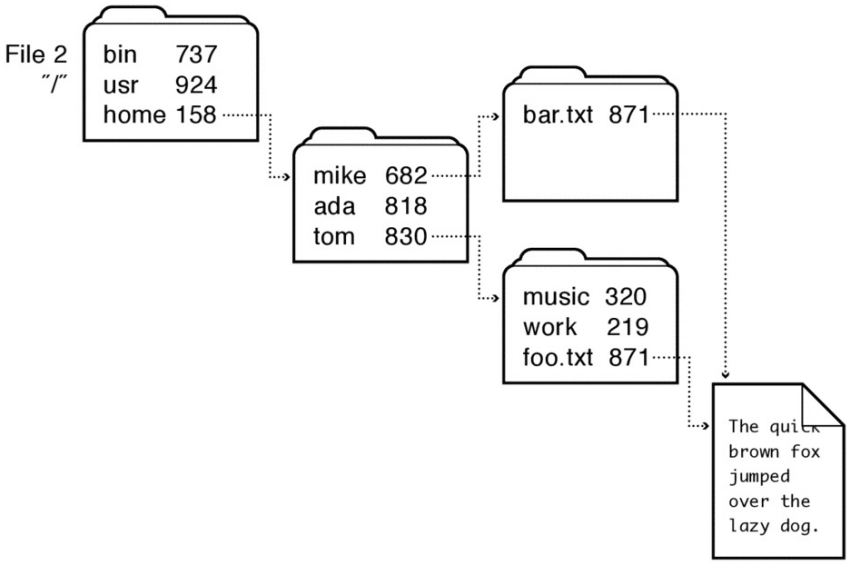


Figure 13.6: Example of a directed acyclic graph directory organization with multiple hard links to a file (figure repeated from Chapter 11).

Hard links are multiple file directory entries that map different path names to the same file number. Because a file number can appear in multiple directories, file systems must ensure that a file is only deleted when the last hard link to it has been removed.

To properly implement garbage collection, file systems use reference counts by storing with each file the number of hard links to it. When a file is created, it has a reference count of one, and each additional hard link made to the file (e.g., link(existingName, newName)) increments its reference count. Conversely, each call to unlink(name) decrements the file's reference count, and when the reference count falls to zero, the underlying file is removed and its resources marked as free.

Rather than mapping a file name to a file number, *soft links* or *symbolic links* are directory entries that map one name to another name.

Figure 13.7: In this directory, the hard links foo.txt and bar.txt and the soft link baz.txt all refer to the same file.

For example, Figure 13.7 shows a directory that contains three names that all refer to the same file. The entries foo.txt and bar.txt are hard links to the same file — number 66212871; baz.txt is a soft link to foo.txt.

Notice that if we remove entry foo.txt from this directory using the unlink system call, the file can still be opened using the name bar.txt, but if we try to open it with the name baz.txt, the attempt will fail.

EXAMPLE: File metadata. Most file systems store a file's metadata (e.g., a file's access time, owner ID, permissions, and size) in a file header that can be found with the file number. One could imagine storing that metadata in a file's directory entry instead. Why is this seldom done?

ANSWER: In file systems that support hard links, storing file metadata in directory entries would be problematic. For example, whenever a file's attribute like its size changed, all of a file's directory entries would have to be located and updated. As another example, if file metadata were stored in directory entries, it would be hard to maintain a file reference count so that the file's resources are freed when and only when the last hard link to the file is removed.

The venerable Microsoft FAT file system stores file metadata in directory entries, but it does not support hard links. □

13.3 Files: Finding Data

Once a file system has translated a file name into a file number using a directory, the file system must be able to find the blocks that belong to that file. In addition to this functional requirement, implementations of files typically target five other goals:

- Support sequential data placement to maximize sequential file access
- Provide efficient random access to any file block
- Limit overheads to be efficient for small files
- Be scalable to support large files
- Provide a place to store per-file metadata such as the file's reference count, owner, access control list, size, and access time

File system designers have a great deal of flexibility to meet these goals. Recall from Section 13.1 that

- A file's *index structure* provides a way to locate each block of the file. Index structures are usually some sort of tree for scalability and to support locality.
- A file system's *free space map* provides a way to allocate free blocks to grow a file. When files grow, choosing which free blocks to use is important for providing good locality. A file system's free space map is therefore often implemented as a bitmap so that it is easy to find a desired number of sequential free blocks near a desired location.
- A file system's *locality heuristics* define how a file system groups data in storage to maximize access performance.

	FAT	FFS	NTFS	ZFS
Index structure	linked list	tree (fixed, assymmetric)	tree (dynamic)	tree (COW, dynamic)
Index structure granularity	block	block	extent	block
Free space management	FAT array	bitmap (fixed)	bitmap in file (file)	space map (log-structured)
Locality heuristics	defrag.	block groups	best fit	write-anywhere
		reserve space	defrag.	block groups

Figure 13.8: Summary of key ideas discussed for four common file systems approaches.

Within this framework, the design space for file systems is large. To understand the trade-offs and to understand the workings of common file systems, we will examine four case study designs that illustrate important implementation techniques and that represent approaches that are in wide use today.

- **FAT.** The Microsoft File Allocation Table (FAT) file system traces its roots to the late 1970s.

  Techniques: The FAT file system uses an extremely simple index structure — a *linked list* — so it is a good place to discuss our discussion of implementation techniques.

  Today: The FAT file system is still widely used in devices like flash memory sticks and digital cameras where simplicity and interoperability are paramount.

- **FFS.** The Unix Fast File System (FFS) was released in the mid 1980s, and it retained many of the data structures in Ritchie and Thompson's original Unix file system from the early 1970s.

  Techniques: FFS uses a *tree-based multi-level index* for its index structure to improve random access efficiency, and it uses a collection of *locality heuristics* to get good spatial locality for a wide range of workloads.

  Today: In Linux, the popular ext2 and ext3 file systems are based on the FFS design.

- **NTFS.** The Microsoft New Technology File System (NTFS) was introduced in the early 1990s as a replacement for the FAT file system.

  Techniques: Like FFS, NTFS uses a tree-based index structure, but the tree is *more flexible than FFS's fixed tree*. Additionally, NTFS optimizes its index structure for sequential file layout by indexing variable-sized *extents* rather than individual blocks.

  Today: NTFS remains the primary file system for Microsoft operating systems such as Windows 7. In addition, the flexible tree and extent techniques are representative of several widely used file systems such as the Linux ext4, XFS, and Reiser4 file systems, the AIX/Linux Journaled File System (JFS), and the Apple Hierarchical File Systems (HFS and HFS+).

- **COW/ZFS.** Copy-on-write (COW) file systems update existing data and metadata blocks by writing new versions to free disk blocks. This approach optimizes write performance: because any data or metadata can be written to any free space on disk, the file system can group otherwise random writes into large, sequential group writes.

  To see how these ideas are implemented, we will examine the open-source ZFS, a prominent copy-on-write file system that was introduced in the early 2000's by Sun Microsystems. ZFS is designed to scale to file systems spanning large numbers of disks, to provide strong data integrity guarantees, and to optimize write performance.

Figure 13.8 summarizes key ideas in these systems that we will detail in the sections that follow.

Sectors vs. pages; blocks vs. clusters; extents vs. runs

Although storage hardware arranges data in *sectors* (for magnetic disk) or *pages* (for flash), file systems often group together a fixed number of disk sectors or flash pages into a larger allocation unit called a *block*. For example, we might format a file system running on a disk with 512 byte sectors to use 4 KB blocks. Aggregating multiple sectors into a block can reduce the overheads of allocating, tracking, and de-allocating blocks,

but it may increase space overheads slightly.

FAT and NTFS refer to blocks as *clusters*, but for consistency we will use the term *block* in our discussions.

Finally, some file systems like NTFS, ext4, and btrfs store data in variable-length arrays of contiguous tracks called *extents* in most file systems and *runs* in NTFS. For consistency, we will use the term *extent* in our discussions.

13.3.1 FAT: Linked List

The Microsoft File Allocation Table (FAT) file system was first implemented in the late 1970s and was the main file system for MS-DOS and early versions of Microsoft Windows. The FAT file system has been enhanced in many ways over the years. Our discussion will focus on the most recent version, FAT-32, which supports volumes with up to 2^{28} blocks and files with up to 2^{32} - 1 bytes.

Index structures. The FAT file system is named for its *file allocation table*, an array of 32-bit entries in a reserved area of the volume. Each file in the system corresponds to a linked list of FAT entries, with each FAT entry containing a pointer to the next FAT entry of the file (or a special "end of file" value). The FAT has one entry for each block in the volume, and the file's blocks are the blocks that correspond to the file's FAT entries: if FAT entry i is the jth FAT entry of a file, then storage block i is the jth data block of the file.

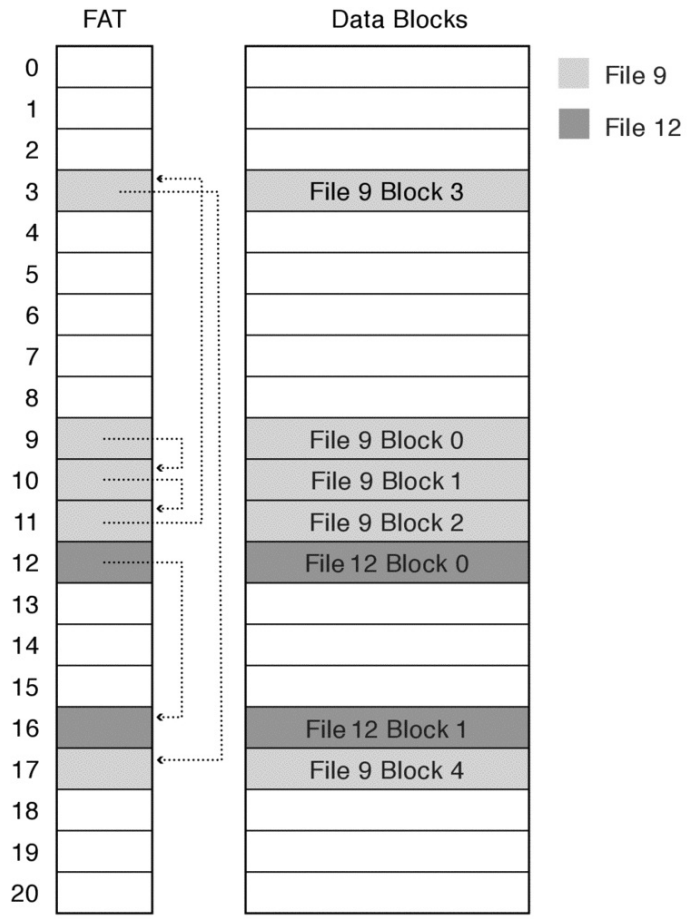


Figure 13.9: A FAT file system with one 5-block file and one 2-block file.

Figure 13.9 illustrates a FAT file system with two files. The first begins at block 9 and contains five blocks. The second begins at block 12 and contains two blocks.

Directories map file names to file numbers, and in the FAT file system, a file's number is the index of the file's first entry in the FAT. Thus, given a file's number, we can find the first FAT entry and block of a file, and given the first FAT entry, we can find the rest of the file's FAT entries and blocks.

Free space tracking. The FAT is also used for free space tracking. If data block i is free, then FAT[i] contains 0. Thus, the file system can find a free block by scanning through the FAT to find a zeroed entry.

Locality heuristics. Different implementations of FAT may use different allocation strategies, but FAT implementations' allocation strategies are usually simple. For example, some implementations use a *next fit* algorithm that scans sequentially through the FAT starting from the last entry that was allocated and that returns the next free entry found.

Simple allocation strategies like this may fragment a file, spreading the file's blocks across the volume rather than achieving the desired sequential layout. To improve performance, users can run a *defragmentation* tool that reads files from their existing locations and rewrites them to new locations with better spatial locality. The FAT defragmenter in Windows XP, for example, attempts to copy the blocks of each file that is spread across multiple extents to a single, sequential extent that holds all the blocks of a file.

Discussion The FAT file system is widely used because it is simple and supported by many operating systems. For example, many flash storage USB keys and camera storage cards use the FAT file system, allowing them to be read and written by almost any computer running almost any modern operating system.

Variations of the FAT file system are even used by applications for organizing data within individual files. For example, Microsoft .doc files produced by versions of Microsoft Word from 1997 to 2007 are actually compound documents with many internal pieces. The .doc format creates a FAT-like file system within the .doc file to manage the objects in the .doc file.

The FAT file system, however, is limited in many ways. For example,

- **Poor locality.** FAT implementations typically use simple allocation strategies such as next fit. These can lead to badly fragmented files.
- **Poor random access.** Random access within a file requires sequentially traversing the file's FAT entries until the desired block is reached.
- **Limited file metadata and access control.** The metadata for each file includes information like the file's name, size, and creation time, but it does not include access control information like the file's owner or group ID, so any user can read or write any file stored in a FAT file system.
- **No support for hard links.** FAT represents each file as a linked list of 32-bit entries in the file allocation table. This representation does not include room for any other file metadata. Instead, file metadata in stored with directory entries with the file's name. This approach demands that each file be accessed via exactly one directory entry, ruling out multiple hard links to a file.
- **Limitations on volume and file size.** FAT table entries are 32 bits, but the top four bits are reserved. Thus, a FAT volume can have at most 2^{28} blocks. With 4 KB blocks, the maximum volume size is limited (e.g., 2^{28} blocks/volume × 2^{12} bytes/block = 2^{40} bytes/volume = 1 TB). Block sizes up to 256 KB are supported, but they risk wasting large amounts of disk space due to internal fragmentation.

  Similarly, file sizes are encoded in 32 bits, so no file can be larger than 2^{32} - 1 bytes (just under 4 GB).
- **Lack of support for modern reliability techniques.** Although we will not discuss reliability until Chapter 14, we note here that FAT does not support the transactional update techniques that modern file systems use to avoid corrupting critical data structures if the computer crashes while writing to storage.

13.3.2 FFS: Fixed Tree

The Unix Fast File System (FFS) illustrates important ideas for both indexing a file's blocks so they can be located quickly and for placing data on disk to get good locality.

In particular, FFS's index structure, called a *multi-level index*, is a carefully structured tree that allows FFS to locate any block of a file and that is efficient for both large and small files.

Given the flexibility provided by FFS's multi-level index, FFS employs two locality heuristics — *block group placement* and *reserve space* — that together usually provide good on-disk layout.

Index structures. To keep track of the data blocks that belong to each file, FFS uses a fixed, asymmetric tree called a *multi-level index*, as illustrated in Figure 13.10.

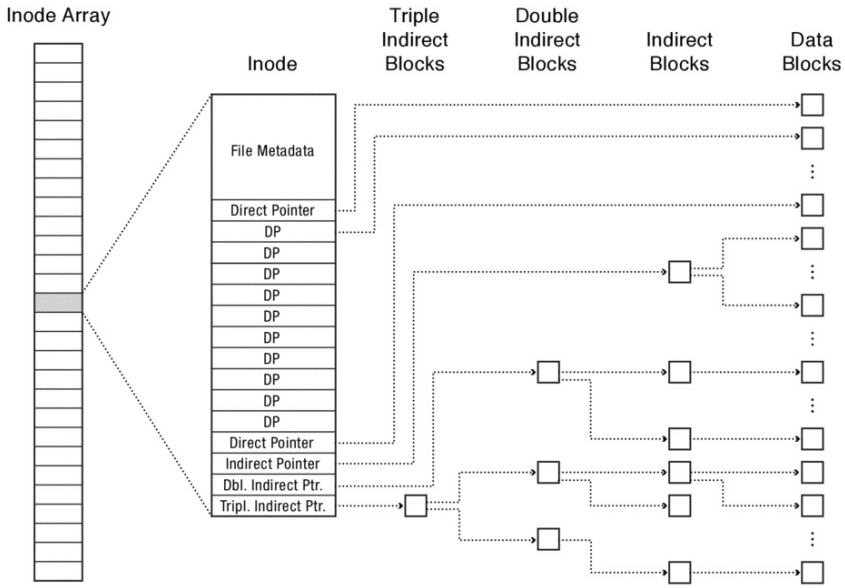


Figure 13.10: An FFS inode is the root of an asymmetric tree whose leaves are the data blocks of a file.

Each file is a tree with fixed-sized data blocks (e.g., 4 KB) as its leaves. Each file's tree is rooted at an *inode* that contains the file's metadata (e.g., the file's owner, access control

permissions, creation time, last modified time, and whether the file is a directory or not).

FFS access control

The FFS inode contains information for controlling access to a file. Access control can be specified for three sets of people:

- **User (owner).** The user that owns the file.
- **Group.** The set of people belonging to a specified Unix group. Each Unix group is specified elsewhere as a group name and list of users in that group.
- **Other.** All other users.

Access control is specified in terms of three types of activities:

- **Read.** Read the regular file or directory.
- **Write.** Modify the regular file or directory.
- **Execute.** Execute the regular file or traverse the directory to access files or subdirectories in it.

Each file's inode stores the identities of the file's user (owner) and group as well as 9 basic access control bits to specify read/write/execute permission for the file's user (owner)/group/other. For example, the command ls -ld / shows the access control information for the root directory:

```
> ls -ld /
drwxr-xr-x  40 root  wheel  1428 Feb  2 13:39 /
```

This means that the file is a directory (d), owned by the user root the group wheel. The root directory can be read, written, and executed (traversed) by the owner (rwx), and it can be read and executed (traversed) but not written by members of group wheel (r-x) and all other users (r-x).

Setuid and setgid programs

In addition to the 9 basic access control bits, the FFS inode stores two important additional bits:

- **Setuid.** When this file is executed by any user (with execute permission) it will be executed with the file owner's (rather than the user's) permission. For example, the lprm program allows a user to remove a job from a printer queue. The print queue is implemented as a directory containing files to be printed, and because we do not want users to be able to remove other users' jobs, this directory is owned by and may only be modified by the root user. So, the lprm program is owned by the root user with the setuid bit set. It can be executed by anyone, but when it runs, it executes with root permissions, allowing it to modify the print queue directory. E.g.,

  ```
  -rwsr-xr-x 1 root root 507674 2010-07-05 12:39 /lusr/bin/lprm*
  ```

  Of course, making a program setuid is potentially dangerous. Here, for example, we rely on the lprm program to verify that actual user is deleting his own print jobs, not someone else's. A bug in the lprm program could let one user remove another's printer jobs. Worse, if the bug allows the attacker to execute malicious code (e.g., via a buffer overflow attack), a bug in lprm could give an attacker total control of the machine. http://www.linuxjournal.com/article/6701
- **Setgid.** The setguid bit is similar to the setuid bit, execept that the file is executed with the file's group permission. For example, on some machines, sendmail executes as a member of group smmsp so that it can access a mail queue file accessible to group smmsp. E.g.,

  ```
  -r-xr-sr-x 1 root smmsp 2264923 2011-06-23 14:51 /lusr/opt/sendmail-8.14.4/lib/mail/sendmail*
  ```

A file's inode (root) also contains array of pointers for locating the file's data blocks (leaves). Some of these pointers point directly to the tree's data leaves and some of them point to internal nodes in the tree. Typically, an inode contains 15 pointers. The first 12 pointers are *direct pointers* that point directly to the first 12 data blocks of a file.

The 13th pointer is an *indirect pointer*, which points to an internal node of the tree called an *indirect block*; an indirect block is a regular block of storage that contains an array of direct pointers. To read the 13th block of a file, you first read the inode to get the indirect pointer, then the indirect block to get the direct pointer, then the data block. With 4 KB blocks and 4-byte block pointers, an indirect block can contain as many as 1024 direct pointers, which allows for files up to a little over 4 MB.

The 14th pointer is a *double indirect pointer*, which points to an internal node of the tree called a *double indirect block*; a double indirect block is an array of indirect pointers, each of which points to an indirect block. With 4 KB blocks and 4-byte block pointers, a double indirect block can contain as many as 1024 indirect pointers. Thus, a double indirect pointer can index as many as $(1024)^2$ data blocks.

Finally, the 15th pointer is a *triple indirect pointer* that points to a *triple indirect block* that contains an array of double indirect pointers. With 4 KB blocks and 4-byte block pointers, a triple indirect pointer can index as many as $(1024)^3$ data blocks containing 4 KB × 1024^3 = $2^{12} \times 2^{30} = 2^{42}$ bytes (4 TB).

All of a file system's inodes are located in an *inode array* that is stored in a fixed location on disk. A file's file number, called an *inumber* in FFS, is an index into the inode array: to

open a file (e.g., foo.txt), we look in the file's directory to find its inumber (e.g., 91854), and then look in the appropriate entry of the inode array (e.g., entry 91854) to find its metadata.

FFS's multi-level index has four important characteristics:

1. **Tree structure.** Each file is represented as a tree, which allows the file system to efficiently find any block of a file.
2. **High degree.** The FFS tree uses internal nodes with many children compared to the binary trees often used for in-memory data structures (i.e., internal nodes have high *degree* or *fan out*). For example, if a file block is 4 KB and a blockID is 4 bytes, then each indirect block can contain pointers to 1024 blocks.

   High degree nodes make sense for on-disk data structures where (1) we want to minimize the number of seeks, (2) the cost of reading several kilobytes of sequential data is not much higher than the cost of reading the first byte, and (3) data must be read and written at least a sector at a time.

   High degree nodes also improve efficiency for sequential reads and writes — once an indirect block is read, hundreds of data blocks can be read before the next indirect block is needed. Runs between reads of double indirect blocks are even larger.
3. **Fixed structure.** The FFS tree has a fixed structure. For a given configuration of FFS, the first set of d pointers always point to the first d blocks of a file; the next pointer is an indirect pointer that points to an indirect block; etc.

   Compared to a dynamic tree that can add layers of indirection above a block as a file grows, the main advantage of the fixed structure is implementation simplicity.
4. **Asymmetric.** To efficiently support both large and small files with a fixed tree structure, FFS's tree structure is asymmetric. Rather than putting each data block at the same depth, FFS stores successive groups of blocks at increasing depth so that small files are stored in a small-depth tree, the bulk of medium files are stored in a medium-depth tree, and the bulk of large files are stored in a larger-depth tree. For example, Figure 13.11 shows a small, 4-block file whose inode includes direct pointers to all of its blocks. Conversely, for the large file shown in Figure 13.10, most of the blocks must be accessed via the triple indirect pointer.

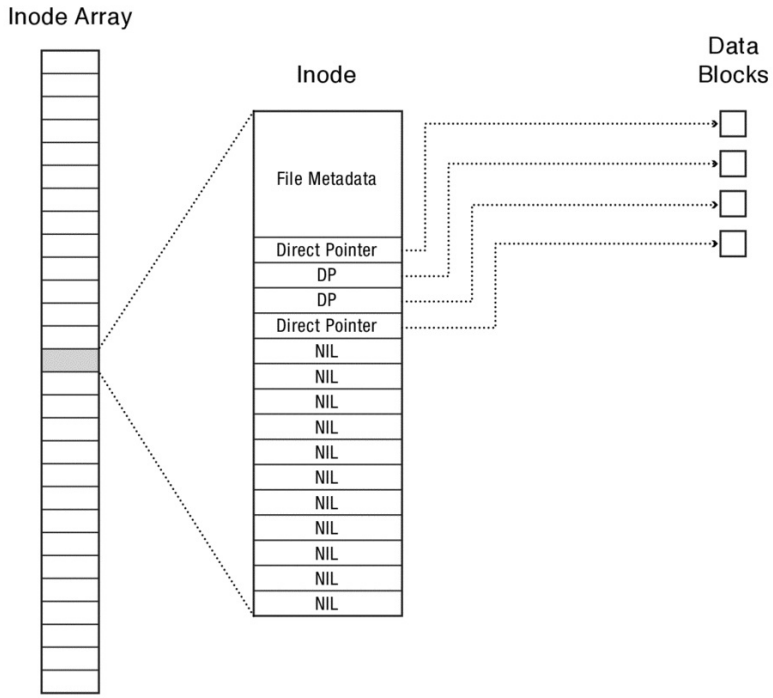


Figure 13.11: A small FFS file whose blocks are all reachable via direct pointers in the inode.

In contrast, if we use a fixed-depth tree and want to support reasonably large files, small files would pay high overheads. With triple indirect pointers and 4 KB blocks, storing a 4 KB file would consume over 16 KB (the 4 KB of data, the small inode, and 3 levels of 4 KB indirect blocks), and reading the file would require reading five blocks to traverse the tree.

The FFS principles are general; many file systems have adopted variations on its basic approach.

EXAMPLE: FFS variation. Suppose BigFS is a variation of FFS that includes in each inode 12 direct, 1 indirect, 1 double indirect, 1 triple indirect, and 1 *quadruple indirect* pointers. Assuming 4 KB blocks and 8-byte pointers, what is the maximum file size this index structure can support?

ANSWER: 12 direct pointers can index 12 × 4 KB = 48 KB.

When used as an internal node, each storage block can contain as many as 4 KB/block / 8 bytes/pointer = 512 pointers/block = 2^9 pointers/block.

So, the indirect pointer points to an indirect block with 2^9 pointers, referencing as much as

2^9 blocks × 2^{12} bytes/block = 2^{21} bytes = 2 MB.

Similarly, the double indirect pointer references as much as $2^9 \times 2^9 \times 2^{12} = 2^{30}$ bytes = 1 GB, the triple indirect pointer references as much as $2^9 \times 2^9 \times 2^9 \times 2^{12} = 2^{39}$ bytes = 512 GB, and the quadruple indirect pointer references as much as $2^9 \times 2^9 \times 2^9 \times 2^9 \times 2^{12} = 2^{48}$ bytes = 256 TB.

So, **BigFS can support files a bit larger than 256.5 TB.** □

Sparse files. Tree-based index structures like FFS's can support *sparse files* in which one or more ranges of empty space are surrounded by file data. The ranges of empty space consume no disk space.

For example, if we create a new file, write 4 KB at offset 0, seek to offset 2^{30}, and write another 4 KB, as Figure 13.12 illustrates, an FFS system with 4 KB blocks will only consume 16 KB — two data blocks, a double indirect block, and a single indirect block.

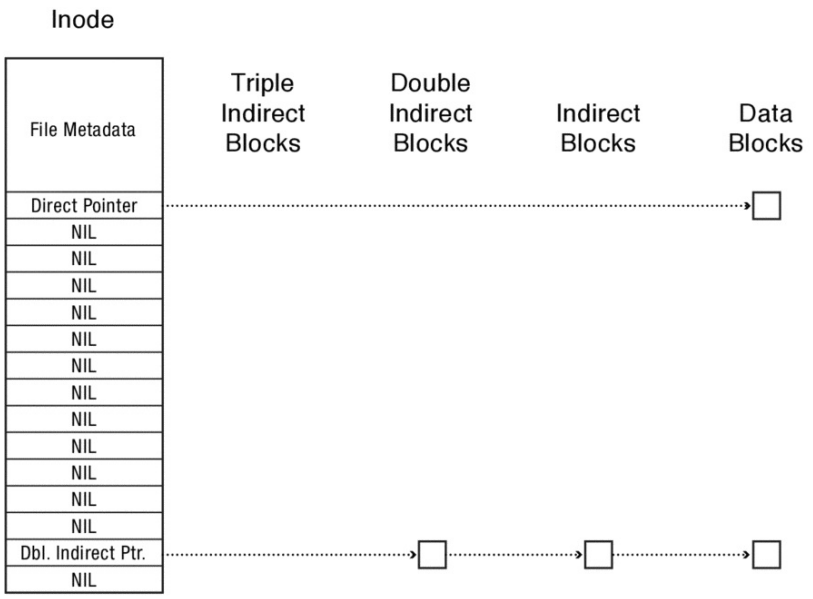


Figure 13.12: A sparse FFS file with two 4 KB bocks, one at offset 0 and one at offset 2^{30}.

In this case, if we list the *size of the file* using the ls command, we see that the file's size is 1.1 GB. But, if we check the *space consumed by the file*, using the du command, we see that it consumes just 16 KB of storage space.

```
>ls -lgGh sparse.dat
-rwx— 1 1.1G 2012-01-31 08:45 sparse.dat*
>du -hs sparse.dat
16K sparse.dat
```

If we read from a hole, the file system produces a zero-filled buffer. If we write to a hole, the file system allocates storage blocks for the data and any required indirect blocks.

Similar to efficient support for sparse virtual memory address spaces, efficient support of sparse files is useful for giving applications maximum flexibility in placing data in a file. For example, a database could store its tables at the start of its file, its indices at 1 GB into the file, its log at 2 GB, and additional metadata at 4 GB.

Sparse files have two important limitations. First, not all file systems support them, so an application that relies on sparse file support may not be portable. Second, not all utilities correctly handle sparse files, which can lead to unexpected consequences. For example, if I read a sparse file from beginning to end and write each byte to a different file, I will observe runs of zero-filled buffers corresponding to holes and write those zero-filled regions to the new file. The result is a new non-sparse file whose space consumption matches its size.

```
>cat sparse.dat > /tmp/notSparse.dat
>ls -lgGh /tmp/notSparse.dat
-rw-r—r— 1 1.1G 2012-01-31 08:54 /tmp/notSparse.dat
 >
>du -hs /tmp/notSparse.dat
1.1G /tmp/notSparse.dat
```

Free space management. FFS's free space management is simple. FFS allocates a *bitmap* with one bit per storage block. The ith bit in the bitmap indicates whether the ith block is free or in use. The position of FFS's bitmap is fixed when the file system is formatted, so it is easy to find the part of the bitmap that identifies free blocks near any location of interest.

Locality heuristics. FFS uses two important locality heuristics to get good performance for many workloads: *block group placement* and *reserved space*.

Block group placement. FFS places data to optimize for the common case where a file's data blocks, a file's data and metadata, and different files from the same directory are accessed together.

Conversely, because everything cannot be near everything, FFS lets different directories' files be far from each other.

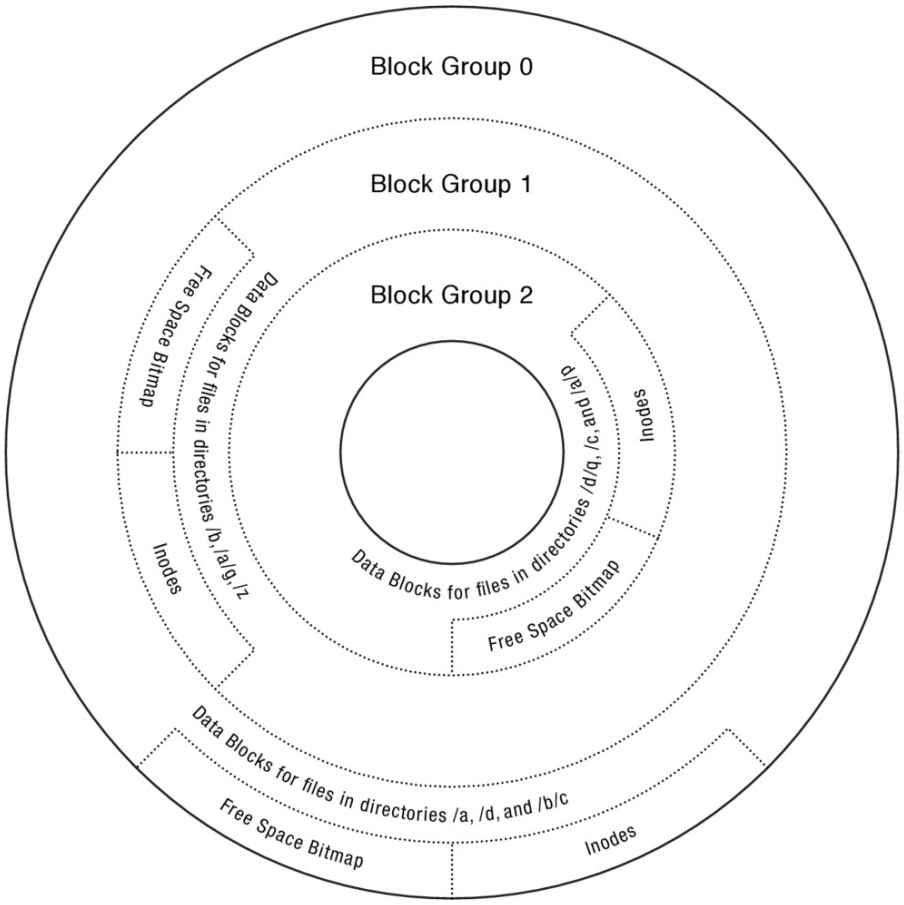


Figure 13.13: FFS divides a disk into block groups, splits free space and inode metadata across block groups, and puts data blocks for the files in a directory in the same block group.

This placement heuristic has four parts:

- **Divide disk into block groups.** As Figure 13.13 illustrates, FFS divides a disk in to sets of nearby tracks called *block groups*. The seek time between any blocks in a block group will be small.
- **Distribute metadata.** Earlier multi-level index file systems put the inode array and free space bitmap in a contiguous region of the disk. In such a centralized metadata arrangement, the disk head must often make seeks between a file's data and its metadata.

  In FFS, the inode array and free space bitmap are still conceptually arrays of records, and FFS still stores each array entry at a well-known, easily calculable location, but the array is now split into pieces distributed across the disk. In particular, each block group holds a portion of these metadata structures as Figure 13.13 illustrates.

  For example, if a disk has 100 block groups, each block group would store 1% of the file system's inodes and the 1% portion of the bitmap that tracks the status of the data blocks in the block group.
- **Place file in block group.** FFS puts a directory and its files in the same block group: when a new file is created, FFS knows the inumber of the new file's directory, and from that it can determine the range of inumbers in the same block group. FFS chooses an inode from that group if one is free; otherwise, FFS gives up locality and selects an inumber from a different block group.

  In contrast with regular files, when FFS creates a new directory, it chooses an inumber from a different block group. Even though we might expect a subdirectory to have some locality with its parent, putting all subdirectories in the same block group would quickly fill it, thwarting our efforts to get locality within a directory.

  Figure 13.13 illustrates how FFS might groups files from different directories into different block groups.
- **Place data blocks.** Within a block group, FFS uses a first-free heuristic. When a new block of a file is written, FFS writes the block to the first free block in the file's block group.

  Although this heuristic may give up locality in the short term, it does so to improve locality in the long term. In the short term, this heuristic might spread a sequence of writes into small holes near the start of a block group rather than concentrating them to a sequence of contiguous free blocks somewhere else. This short term sacrifice brings long term benefits, however: fragmentation is reduced, the block will tend to have a long run of free space at its end, subsequent writes are more likely to be sequential.

  The intuition is that a given block group will usually have a handful of holes scattered through blocks near the start of the group and a long run of free space at the end of the group. Then, if a new, small file is created, its blocks will likely go to a few of the small holes, which is not ideal, but which is acceptable for a small file. Conversely, if a large file is created and written from beginning to end, it will tend to have the first few blocks scattered through the holes in the early part of the block, but then have the bulk of its data written sequentially at the end of the block group.

  If a block group runs out of free blocks, FFS selects another block group and allocates blocks there using the same heuristic.

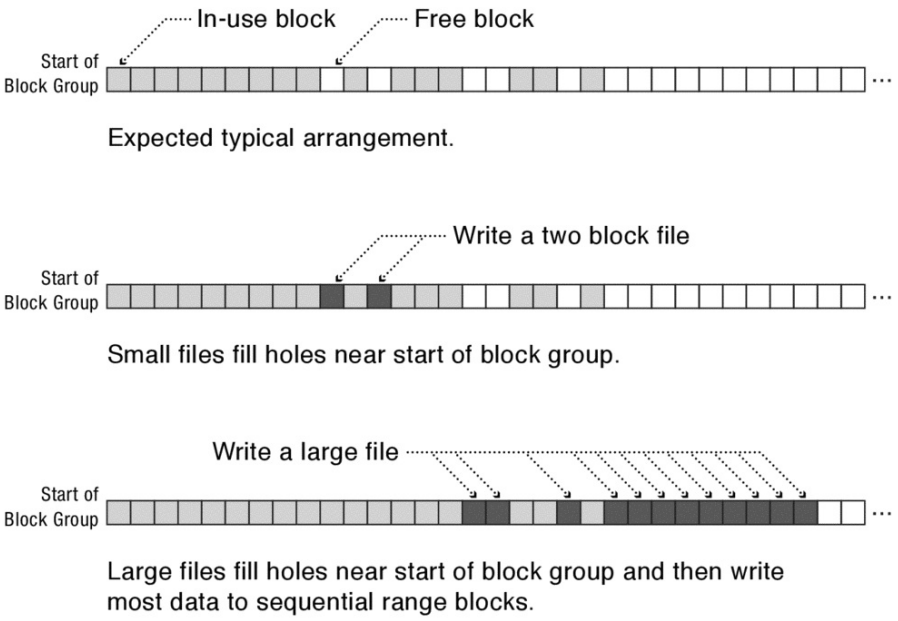


Figure 13.14: FFS's block placement heuristic is to put each new file block in the first free block in that file's block group

Reserved space. Although the block group heuristic can be effective, it relies on there being a significant amount of free space on disk. In particular, when a disk is nearly full, there is little opportunity for the file system to optimize locality. For example, if a disk has only a few kilobytes of free sectors, most block groups will be full, and others will have only a few free blocks; new writes will have to be scattered more or less randomly around the disk.

FFS therefore reserves some fraction of the disk's space (e.g., 10%) and presents a slightly reduced disk size to applications. If the actual free space on the disk falls below the reserve fraction, FFS treats the disk as full. For example, if a user's application attempts to write a new block in a file when all but the reserve space is consumed, that write will fail. When all but the reserve space is full, the super user's processes will still be able to allocate new blocks, which is useful for allowing an administrator to log in and clean things up.

The reserved space approach works well given disk technology trends. It sacrifices a small amount of disk capacity, a hardware resource that has been improving rapidly over recent decades, to reduce seek times, a hardware property that is improving only slowly.

13.3.3 NTFS: Flexible Tree With Extents

The Microsoft New Technology File System (NTFS), released in 1993, improved on Microsoft's FAT file system with many new features including new index structures to improve performance, more flexible file metadata, improved security, and improved reliability.

We will discuss some of NTFS's reliability features in Chapter 14. Here, we will focus on how NTFS stores data and metadata.

Index structures. Whereas FFS tracks file blocks with a fixed tree, NTFS and many other recent file systems such as Linux ext4 and btrfs track *extents* with *flexible trees*.

- **Extents.** Rather than tracking individual file blocks, NTFS tracks *extents*, variable-sized regions of files that are each stored in a contiguous region on the storage device.
- **Flexible tree and master file table.** Each file in NTFS is represented by a variable-depth tree. The extent pointers for a file with a small number of extents can be stored in a shallow tree, even if the file, itself, is large. Deeper trees are only needed if the file becomes badly fragmented.

The roots of these trees are stored in a master file table that is similar to FFS's inode array. NTFS's *master file table (MFT)* stores an array of 1 KB MFT records, each of which stores a sequence of variable-size *attribute records*. NTFS uses attribute records to store both data and metadata — both are just considered attributes of a file.

Some attributes can be too large to fit in an MFT record (e.g., a data extent) while some can be small enough to fit (e.g., a file's last modified time). An attribute can therefore be *resident* or *non-resident*. A *resident attribute* stores its contents directly in the MFT record while a *non-resident attribute* stores extent pointers in its MFT record and stores its contents in those extents.

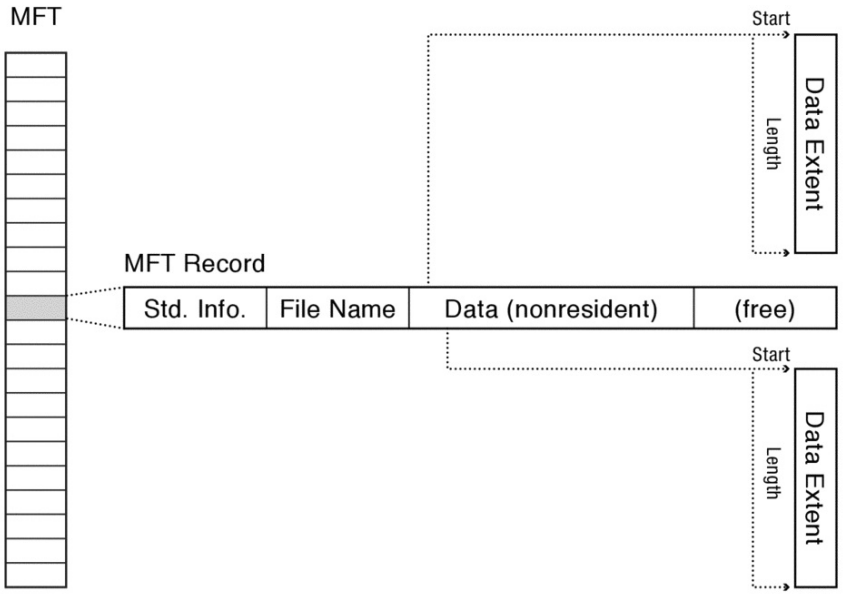


Figure 13.15: NTFS index structures and data for a basic file with two data extents.

Figure 13.15 illustrates the index structures for a basic NTFS file. Here, the file's MFT record includes a *non-resident data* attribute, which is a sequence of extent pointers, each of which specifies the starting block and length in blocks of an extent of data. Because extents can hold variable numbers of blocks, even a multi-gigabyte file can be represented by one or a few extent pointers in an MFT record, assuming file system fragmentation is kept under control.

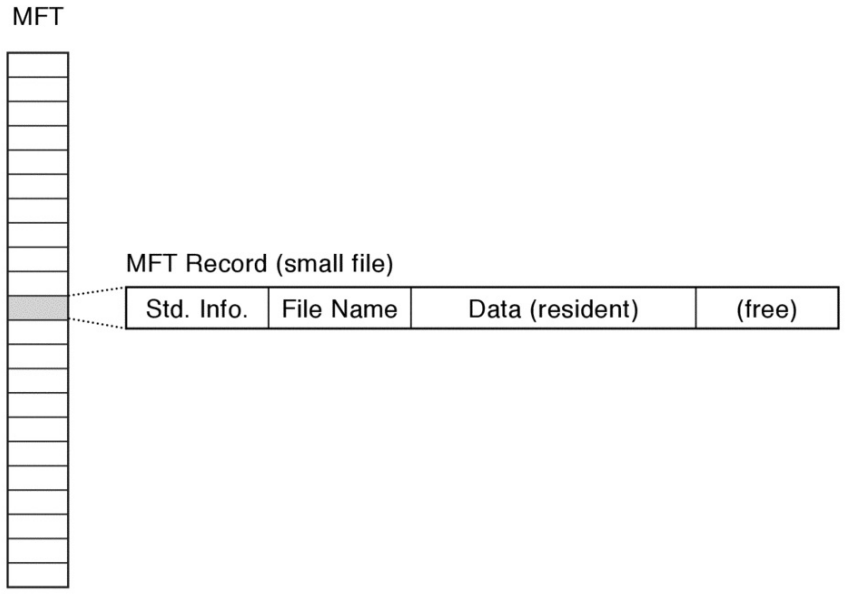


Figure 13.16: A small file's data can be *resident*, meaning that the file's data is stored in its MFT record.

If a file is small, the data attribute may be used to store the file's actual contents right in its MFT record as a resident attribute as Figure 13.16 illustrates.

An MFT record has a flexible format that can include range of different attributes. In addition to data attributes, three common metadata attribute types include:

- **Standard information.** This attribute holds standard information needed for all files. Fields include the file's creation time, modification time, access time, owner ID, and security specifier. Also included is a set of flags indicating basic information like whether the file is a read only file, a hidden file, or a system file.
- **File name.** This attribute holds the file's name and the file number of its parent directory. Because a file can have multiple names (e.g., if there are multiple hard links to the file), it may have multiple file name attributes in its MFT record.
- **Attribute list.** Because a file's metadata may include a variable number of variable sized attributes, a file's metadata may be larger than a single MFT record can hold. When this case occurs, NTFS stores the attributes in multiple MFT records and includes an attribute list in the first record. When present, the attribute list indicates which attributes are stored in which MFT records. For example, Figure 13.17 shows MFT records for two files, one whose attributes are contained in a single MFT record and one of whose attributes spans two MFT records.

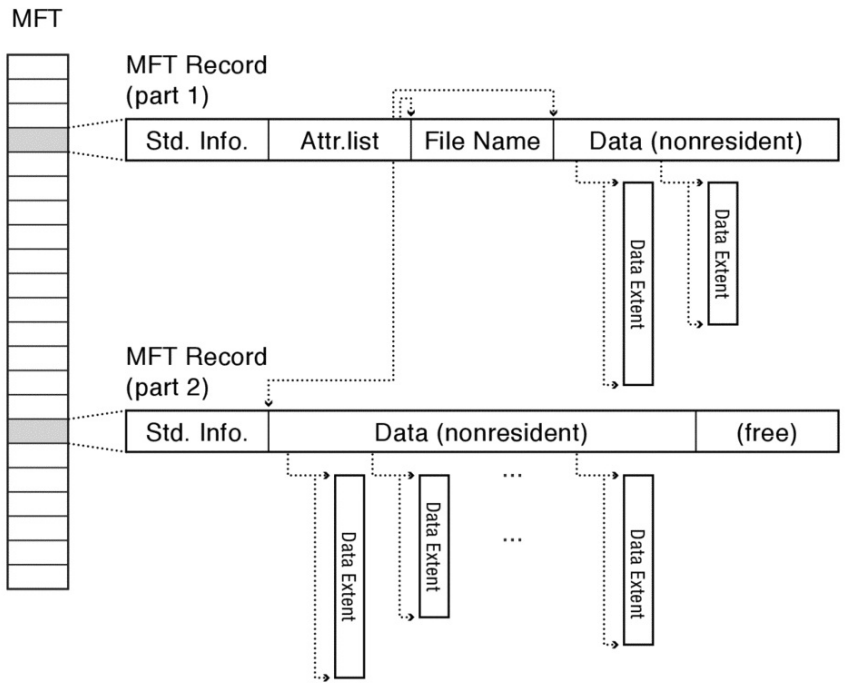


Figure 13.17: Most NTFS files store their attributes in a single MFT record, but a file's attributes can grow to span multiple MFT records. In those cases, the first MFT record includes an *attribute list* attribute that indicates where to find each attribute record.

As Figure 13.18 illustrates, a file can go through four stages of growth, depending on its size and fragmentation. First, a small file may have its contents included in the MFT record as a resident data attribute. Second, more typically, a file's data lies in a small number of extents tracked by a single non-resident data attribute. Third, occasionally if a file is large and the file system fragmented, a file can have so many extents that the extent pointers will not fit in a single MFT record. In this case, as a file can have multiple non-resident data attributes in multiple MFT records, with the attribute list in the first MFT record indicating which MFT records track which ranges of extents. Fourth and finally, if a file is huge or the file system fragmentation is extreme, a file's attribute list can be made non-resident, allowing almost arbitrarily large numbers of MFT records.

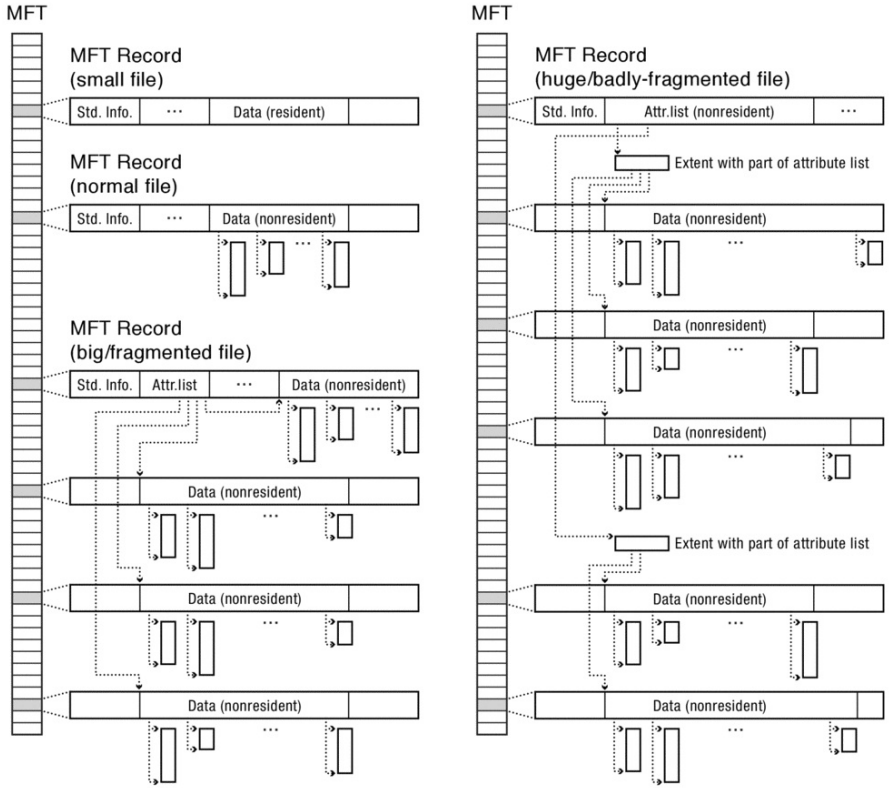


Figure 13.18: An NTFS file's data attribute can be in: (i) a *resident data attribute*, (ii) extents tracked by a *single non-resident data attribute*, (iii) extents tracked by *multiple non-resident data attributes* in multiple MFT entries tracked by a *resident attribute list attribute*, or (iv) extents tracked by *multiple non-resident data attributes* stored in multiple MFT entries tracked by a *non-resident attribute list attribute*.

Metadata files. Rather than doing ad-hoc allocation of special regions of disk for file system metadata like free space bitmaps, NTFS stores almost all of its metadata in about a dozen ordinary files with well-known low-numbered file numbers. For example, file number 5 is the root directory, file number 6 is the free space bitmap, and file number 8 contains a list of the volume's bad blocks.

File number 9, called $Secure, contains security and access control information. NTFS has a flexible security model in which a file can be associated with a list of users and groups, with specific access control settings for each listed principal. In early versions of NTFS, such an access control list was stored with each file, but these lists consumed a nontrivial amount of space and many lists had identical contents. So, current implementations of NTFS store each unique access control list once in the special $Secure file, indexed by a fixed-length unique key. Each individual file just stores the appropriate fixed-length key in its MFT record, and NTFS uses a file's security key to find the appropriate access control

list in the $Secure file.

Even the master file table, itself, is stored as a file, file number 0, called $MFT. So, we need to find the first entry of the MFT in order to read the MFT! To locate the MFT, the first sector of an NTFS volume includes a pointer to the first entry of the MFT.

Storing the MFT in a file avoids the need to statically allocate all MFT entries as a fixed array in a predetermined location. Instead, NTFS starts with a small MFT and grows it as new files are created and new entries are needed.

Locality heuristics.

Most implementations of NTFS use a variation of *best fit*, where the system tries to place a newly allocated file in the smallest free region that is large enough to hold it. In NTFS's variation, rather than trying to keep the allocation bitmap for the entire disk in memory, the system caches the allocation status for a smaller region of the disk and searches that region first. If the bitmap cache holds information for areas where writes recently occurred, then writes that occur together in time will tend to be clustered together.

An important NTFS feature for optimizing its best fit placement is the SetEndOfFile() interface, which allows an application to specify the expected size of a file at creation time. In contrast, FFS allocates file blocks as they are written, without knowing how large the file will eventually grow.

To avoid having the master file table file ($MFT) become fragmented, NTFS reserves part of the start of the volume (e.g., the first 12.5% of the volume) for MFT expansion. NTFS does not place file blocks in the MFT reserve area until the non-reserved area is full, at which point it halves the size of the MFT reserve area and continues. As the volume continues to fill, NTFS continues to halve the reserve area until it reaches the point where the remaining reserve area is more than half full.

Finally, Microsoft operating systems with NTFS include a defragmentation utility that takes fragmented files and rewrites them to contiguous regions of disk.

13.3.4 Copy-On-Write File Systems

When updating an existing file, *copy-on-write (COW) file systems* do not overwrite the existing data or metadata; instead, they write new versions to new locations.

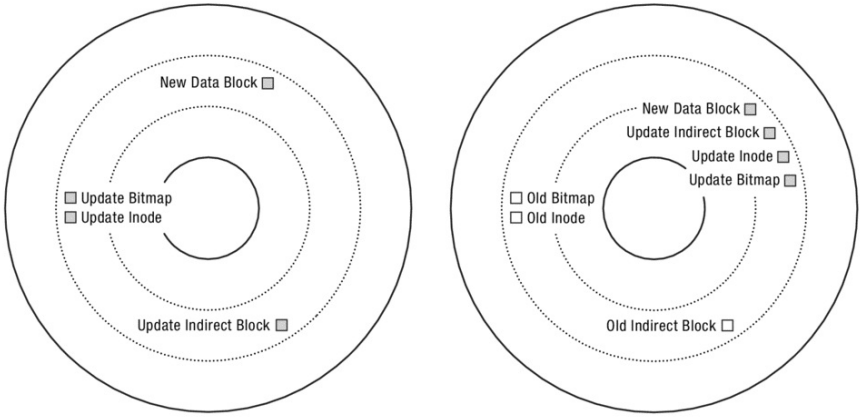


Figure 13.19: An update-in-place file system (left) updates data and metadata in their existing locations, while a copy-on-write file system (right) makes new copies of data and metadata whenever they are updated.

COW file systems do this to optimize writes by transforming random I/O updates to sequential ones. For example, when appending a block to a file, a traditional, update-in-place file system might seek to and update its free space bitmap, the file's inode in the inode array, the file's indirect block, and the file's data block. In contrast, a COW file system might just find a sequential run of free space and write the new bitmap, inode, indirect block, and data block there as illustrated in Figure 13.19.

Several technology trends are driving widespread adoption of COW file systems:

- **Small writes are expensive.** Disk performance for large sequential writes is much better than for small random writes. This gap is likely to continue to grow because bandwidth generally improves faster than seek time or rotational latency: increasing storage density can increase bandwidth even if the rotational speed does not increase.

  As a result the benefits of converting small random writes to large sequential ones is large and getting larger.

- **Small writes are especially expensive on RAID.** Redundant arrays of inexpensive disks (RAIDs) are often used to improve storage reliability. However, as we will discuss in the next chapter, updating a single block stored with parity on a RAID requires four disk I/Os: we must read the old data, read the old parity, write the new data, and write the new parity. In contrast, RAIDs are efficient when an entire stripe — all of the blocks sharing the same parity block — are updated at once. In that case, no reads are needed, each new data block is written, and the parity update is amortized across the data blocks in the stripe.

  Widespread use of RAIDs magnifies the benefits of converting random writes to sequential ones.

- **Caches filter reads.** For many workloads, large DRAM caches can handle essentially all file system reads. But our ability to use DRAM to buffer writes is limited by the need to durably store data soon after it is written.

  Thus, the cost of writes often dominates performance, so techniques that optimize write performance are appealing.

- **Widespread adoption of flash storage.** Flash storage has two properties that make the COW techniques important. First, in order to write a small (e.g., 4 KB) flash page, one must first clear the large (e.g., 512 KB) erasure block on which it resides. Second, each flash storage element can handle a limited number of write-erase cycles before wearing out, so *wear leveling* — spreading writes evenly across all cells — is important for maximizing flash endurance.

  A flash drive's flash translation layer uses COW techniques to virtualize block addresses, allowing it to present a standard interface to read and write specific logical pages while internally redirecting writes to pages on already-cleared erasure blocks and while moving existing data to new physical pages so that their current erasure blocks can be cleared for future writes.

  Note that a flash drive's flash translation layer operates below the file system — standard update in place or COW file systems are still used over that layer. But, flash translation layers are constructed using the same basic principles as the COW file systems discussed here.

- **Growing capacity enables versioning.** Large storage capacities make it attractive for file systems to provide interfaces by which users can access old versions of their files.

  Since updates in COW systems do not overwrite old data with new, supporting versioning is relatively straightforward, as we discuss below.

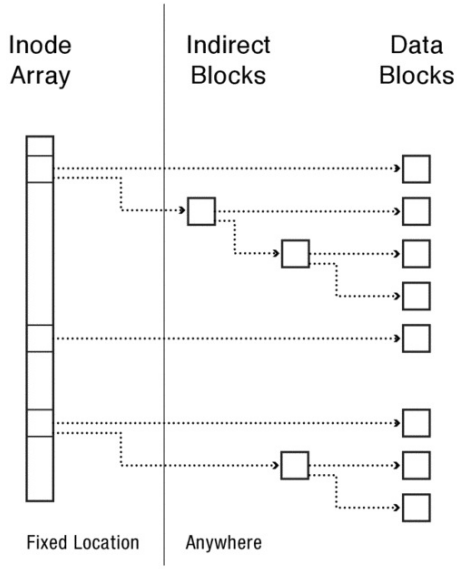


Figure 13.20: A traditional, update-in-place file system, such as FFS.

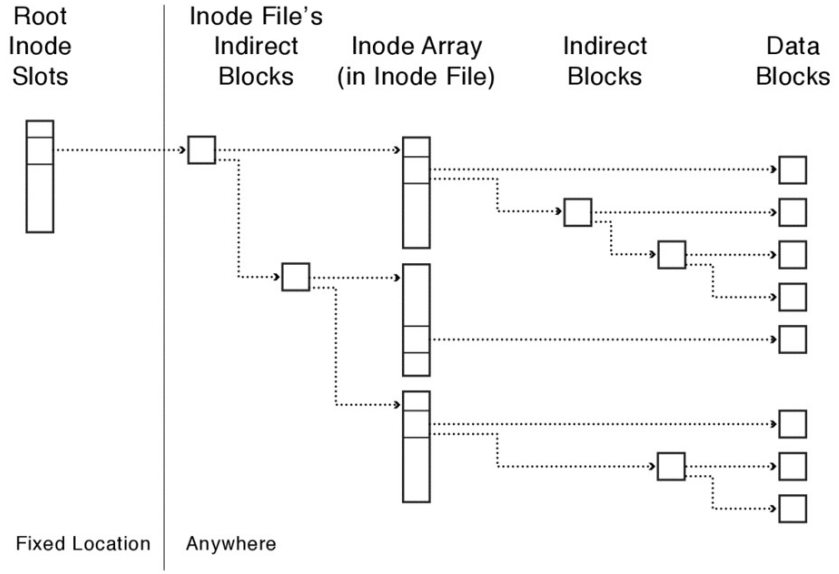


Figure 13.21: A simple copy-on-write (COW) file system.

Implementation principles. Figures 13.20 and 13.21 illustrate the core idea of COW file systems by comparing a traditional file system (FFS in this case) with a COW implementation that uses the same basic index structures.

In the traditional system (Figure 13.20), a file's indirect nodes and data blocks can be located anywhere on disk, and given a file's inumber, we can find its inode in a fixed location on disk.

In the COW version (Figure 13.21), we do not want to overwrite inodes in place, so we must make them mobile. A simple way to do that is to store them in a file rather than in a fixed array. Of course, that is not quite the end of the story — we still need to be able to find the inode file's inode, called the *root inode*.

The simplest thing to do would be to store the root inode in a fixed location. If we did that, then we could find any file's inode by using the root inode to read from a computed offset in the inode file, and from that we could find its blocks.

However, it is useful to make even the root inode copy-on-write. For example, we do not want to risk losing the root inode if there is a crash while it is being written. A solution is to include a monotonically increasing version number and a checksum in the root inode and to keep a small array of slots for the current and recent root inodes, updating the oldest one when a write occurs. After a crash, we scan all of the slots to identify the newest root inode that has a correct checksum.

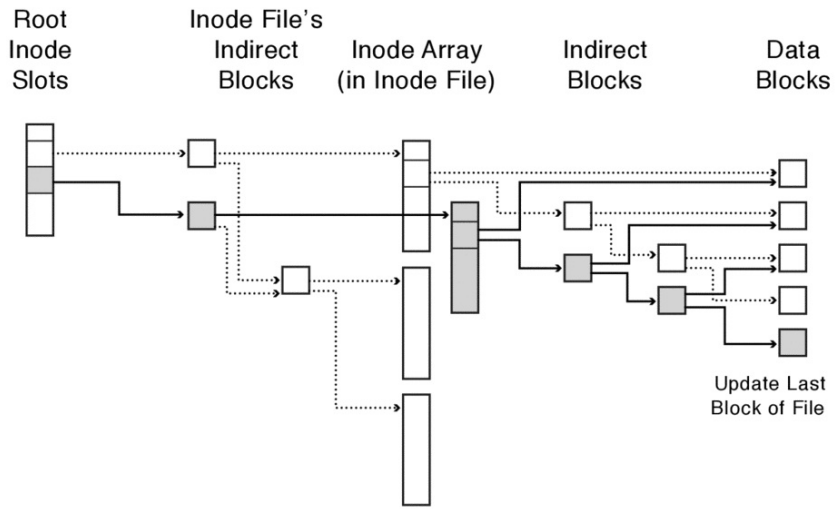


Figure 13.22: In a COW file system, writing a data block causes the system to allocate new blocks for and to write the data block and all nodes on the path from that data block to the root inode.

In this design, all the file system's contents are stored in a tree rooted in the root inode, when we update a block, we write it — and all of the blocks on the path from it to the root — to new locations. For example, Figure 13.22 shows what happens when one block of a file is updated in our simple COW system.

ZFS index structures. To better understand how copy-on-write file systems are implemented, we will look at the open source ZFS file system.

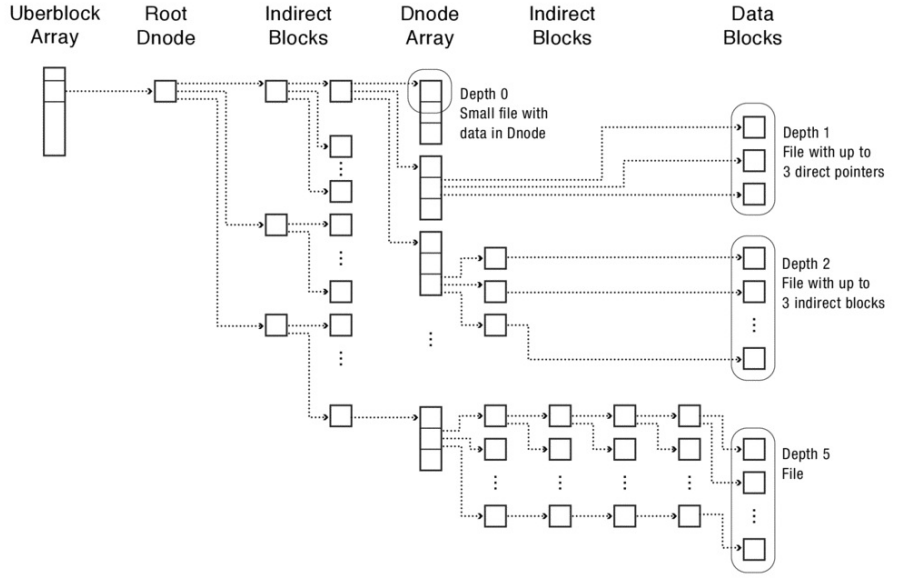


Figure 13.23: ZFS index structures. Note that this diagram is slightly simplified. In reality, there are a few more levels of indirection between the uberblock and a file system's root dnode.

As Figure 13.23 illustrates, the root of a ZFS storage system is called the *uberblock*. ZFS keeps an array of 256 uberblocks in a fixed storage location and rotates successive versions among them. When restarting, ZFS scans the uberblock array and uses the one with a valid checksum that has the highest sequence number.

The current uberblock conceptually includes a pointer to the current root dnode, which holds the dnode array for a ZFS file system. (We say "conceptually" because we are simplifying things a bit here. Once you have read this description, see the sidebar if you want the gory details.)

The basic metadata object in ZFS is called a *dnode*, and it plays a role similar to an inode in FFS or an MFT entry in NTFS: a file is represented by variable-depth tree whose root is a dnode and whose leaves are its data blocks. A dnode has space for three block pointers, and it has a field that specifies the tree's depth: zero indicates that the dnode stores the file's data; one means that the pointers are direct pointers to data blocks; two means that

the dnode's pointers point to indirect blocks, which point to data blocks; three means that the dnode's pointers point to double indirect blocks; and so on, up to six levels of indirection.

Data block and indirect block sizes are variable from 512 bytes to 128KB and specified in a file's dnode. Note, however, that even a 128 KB indirect node holds fewer block pointers than you might expect because each block pointer is a 128 byte structure.

ZFS's block pointers are relatively large structures because they include fields to support advanced features like large storage devices, block compression, placing copies of the same block on different storage devices, file system snapshots, and block checksums. Fortunately, we can ignore these details and just treat each block pointer structure as a (rather large) pointer.

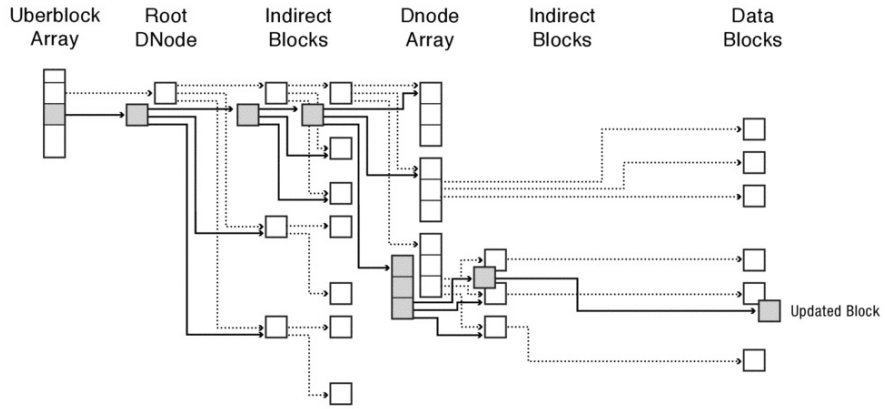


Figure 13.24: Updating a block of a ZFS file

Figure 13.24 shows what happens when we update the last block in a 2-level ZFS file. We allocate a new data block and store the new data in it, but that means that we need to update the indirect block that points to it. So, we allocate a new indirect block and store the version with the updated pointer there, but that means we need to update the indirect pointer that points to it. And so on, up through the file's dnode, the indirect blocks that track the dnode array, the root dnode, and the uberblock.

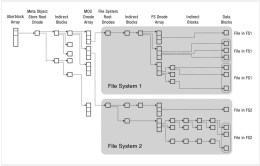
ZFS uberblock, meta-root dnode, and root dnodes

For simplicity, the body of the text describes the uberblock as pointing directly to the file system's root dnode.

In reality, there are a few additional levels of indirection to allow multiple file systems and snapshots to share a ZFS storage pool under a single uberblock. The uberblock has a pointer to a meta-root dnode (called the Meta Object Store dnode in ZFS terminology).

The meta-root dnode tracks the meta-root dnode array. The meta-root dnode array is used by what is essentially a little file system with hierarchical directories that provide mappings from the names of file systems to "files" that store the metadata for each file system, including a pointer to the block where the file system's root dnode is (currently) stored.

So, a more complete picture looks like this:



ZFS space map. ZFS's space maps track free space in a way designed to scale to extremely large storage systems.

One concern the ZFS designers had with bitmaps was that the size of a bitmap grows linearly with storage capacity and can become quite large for large-scale systems. For example, with 4 KB block size, a file server with 1 PB of disk space would have 32 GB of bitmaps.

Large bitmaps affect both a server's memory requirements and the time needed to read the bitmaps on startup. Although one might attempt to cache a subset of the bitmap in memory and only allocate from the currently cached subset, we cannot control when blocks are freed. For workloads in which frees have poor locality, caching will be ineffective.

ZFS's space maps use three key ideas to scale to large storage systems:

- **Per block group space maps.** ZFS maintains a space map for each block group, it restricts allocation of new blocks to a subset of block groups at any given time, and it keeps those block groups' space maps in memory.
- **Tree of extents.** Each block group's free space is represented as an AVL tree of extents. The tree allows ZFS to efficiently find a free extent of a desired size, and its search performance does not degrade as the block group becomes full.
- **Log-structured updates.** As noted above, caching a portion of a space map works for allocations but it may not help frees. Therefore, rather than directly updating the on-disk spacemap, ZFS simply appends spacemap updates to a log. When a block group is activated for allocation, ZFS reads in the most recently stored spacemap and then it reads all subsequently logged updates to bring the space map up to date. After applying updates to the in-memory spacemap, ZFS can store the new spacemap to limit the length of its update log.

ZFS locality heuristics. We started the discussion of COW file systems by saying that they are designed to optimize write performance, but the example in Figure 13.24 make it sound like ZFS does a lot of work just to update a block. ZFS does two important things to optimize write behavior:

- **Sequential writes.** Because almost everything in ZFS is mobile, almost all of these updates can be grouped into a single write to a free range of sequential blocks on disk. Only the uberblock needs to go elsewhere. Because sequential writes are much faster than random ones, ZFS and other COW file systems can have excellent write performance even though they write more metadata than update-in-place file systems.
- **Batched updates.** Figure 13.24 shows what happens when we update a single block of a single file, but ZFS does not typically write one update at a time. Instead, ZFS updates several seconds of updates and writes them to disk as a batch. So, updates to a file's dnode and indirect nodes may be amortized over many writes to the file, and updates to the uberblock, root dnode, and the dnode array's indirect blocks may be amortized over writes to many files.

When it is time to write a batch of writes, ZFS needs to decide where to write the new block versions. It proceeds in three steps:

- **Choose a device.** A ZFS storage pool may span multiple devices, so the first step is to choose which device to use. To maximize throughput by spreading load across devices, ZFS uses a variation of round robin with two tweaks. First, to even out space utilization, ZFS biases selection towards devices with large amounts of free space. Second, to maintain good locality for future reads, ZFS places about 512 KB on one device before moving on to the next one.
- **Choose a block group.** ZFS divides each device into several hundred groups of sequential blocks. ZFS's first choice for is to continue to use the block group it used most recently. However, if that group is so full or fragmented that its largest free region is smaller than 128 KB, ZFS selects a new block group.

  New block group selection is biased towards groups that have more free space, that are nearer the outer edge of a disk (to improve bandwidth), and that have recently been used to store some data (to limit the range of tracks across which the disk head must seek).
- **Choose a block within the group.** To maximize opportunities to group writes together, ZFS uses first fit allocation within a block group until the group is nearly full. At that point, it falls back on best fit to maximize space utilization.

Partitioning, Formatting, and Superblocks

How does an operating system know where to find FFS's inode array, NTFS's MFT, or ZFS's uberblock? How does it know how large these structures are? How does it even know what type of file system is on a disk?

A disk device's space can be divided into multiple *partitions*, each of which appears a separate (smaller) virtual storage device that can be formatted as a separate file system. To partition a disk, an operating system writes a special record (e.g., a master boot record (MBR) or GUID partition table (GPT)) in the first blocks of the disk. This record includes the disk's unique ID, size, and the list of the disk's partitions. Each partition record stores the partition's type (e.g., general file system partition, swap partition, RAID partition, bootable partition), partition ID, partition name, and the partition's starting and ending blocks.

To improve reliability, operating systems store multiple copies of a disk's partition table — often in the first few and last few of a disk's blocks.

Once a disk has been partitioned, the operating system can *format* some or all of the partitions by initializing the partition's blocks according to the requirements of the type of file system being created.

Formatting a file system includes writing a *superblock* that identifies the file system's type and its key parameters such as its type, block size, and inode array or MBR location and size. Again, for reliability, a file system typically stores multiple copies of its superblock at several predefined locations.

Then, when an operating system boots, it can examine a disk to find its partitions, and it can examine each partition to identify and configure its file systems.

13.4 Putting It All Together: File and Directory Access

In Section 13.2 we saw that directories are implemented as files, containing file name to file number mappings, and in Section 13.3 we saw that files are implemented using an index structure — typically a tree of some sort — to track the file's blocks.

In this section, we walk through the steps FFS takes to read a file, given that file's name. The steps for the other file systems we have discussed are similar.

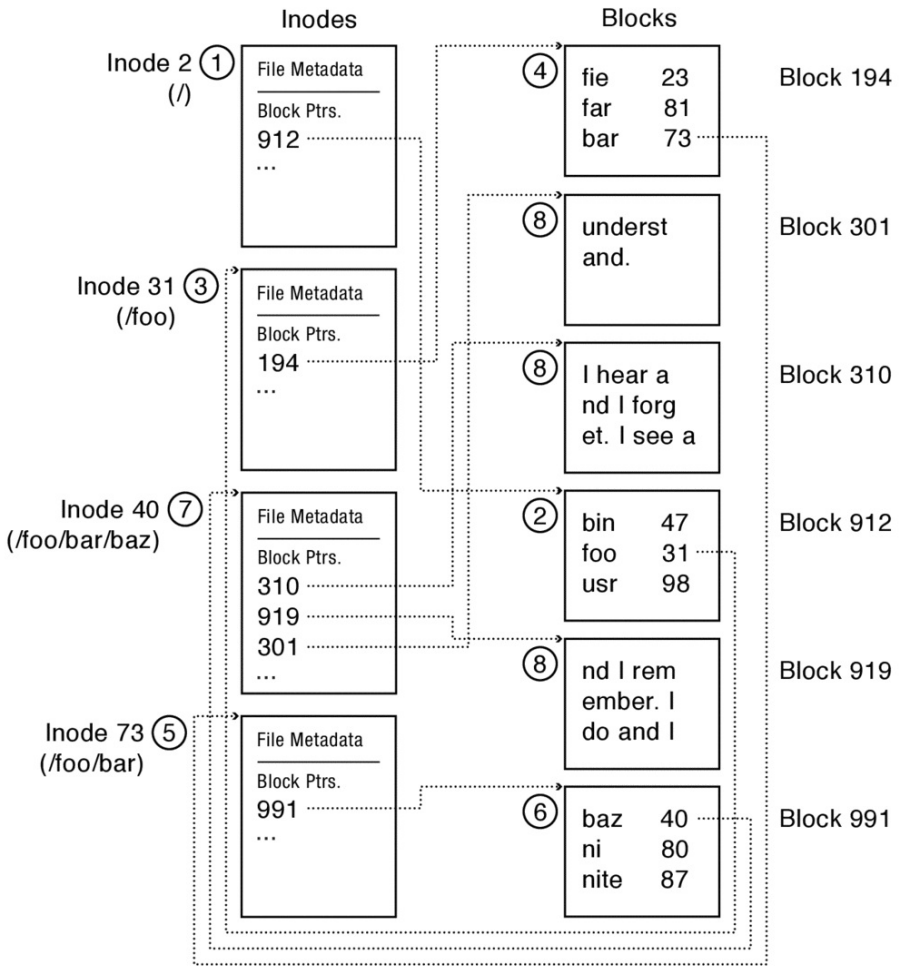


Figure 13.25: The circled numbers identify the steps required to read /foo/bar/baz in the FFS file system.

Suppose we want to read the file /foo/bar/baz.

First, we must read the root directory / to determine /foo's inumber. Since we already know the root directory's inumber (it is a pre-agreed number compiled into the kernel, e.g., 2), we open and read file 2's inode in step 1 in Figure 13.25. Recall that FFS stores pieces of the inode array at fixed locations on disk, so given a file's inumber it is easy to find and read the file's inode.

From the root directory's inode, we extract the direct and indirect block pointers to determine which block stores the contents of the root directory (e.g., block 48912 in this example). We can then read that block of data to get the list of name to inumber mappings in the root directory and discover that directory file /foo has inumber 231 (step 2).

Now that we know /foo's inumber, in step 3 we can read inode 231 to find where /foo's data blocks are stored — block 1094 in the example. We can then read those blocks of data to get the list of name to inumber mappings in the /foo directory and discover that directory file /foo/bar has inumber 731 (step 4).

We follow similar steps to read /foo/bar's inode (step 5) and data block 30991 (step 6) to determine /foo/bar/baz inumber 402.

Finally, in step 7, we read /foo/bar/baz's inode (402), and in step 8, we read its data blocks (89310, 14919, and 23301): "I hear and I forget. I see and I remember. I do and I understand."

This may seem like a lot of steps just to read a file. Most of the time, we expect much of this information to be cached so that some steps can be avoided. For example, if the inodes and blocks for / and /foo are cached, then we would skip steps 1 to 4. Also, once file /foo/bar/baz has been opened, the open file data structure in the operating system will include the file's inumber so that individual reads and writes of the file can skip steps 1 to 6 (and step 7 while the inode is cached).

EXAMPLE: Reading a file. What would you get if you read the file /foo/fie in the FFS file system illustrated in Figure 13.25?

ANSWER: First we read the root inode (inode 2) and file (block 48912), then /foo's inode (inode 231) and file (block 1094), and then /foo/fie's inode (inode 402 again) and file (blocks 89310, 14919, and 23301 again) — /foo/bar/baz and /foo/fie are hard links to the same file. □

13.5 Summary and Future Directions

We are seeing significant shifts in the technologies and workloads that drive file system design.

Practical solid state storage technologies like flash memory change the constraints around which file systems can be designed. Random access performance that is good both in relative terms compared to sequential access performance and in absolute terms provide opportunities to reconsider many aspects of file system design — directories, file metadata structures, block placement — that have been shaped by the limitations of magnetic disks. Promising future solid state storage technologies like phase change memory or memristors may allow even more dramatic restructuring of file systems to take advantage of their even better performance and their support for fine-grained writes of a few bytes or words.

On the other hand, the limited lifetime and capacity of many solid state technologies may impose new constraints on file system designs. Perhaps we should consider hybrid file systems that, for example, store metadata and the content of small files in solid state storage and the contents of large files on magnetic disks.

Even the venerable spinning disk continues to evolve rapidly, with capacity increases continuing to significantly outpace performance improvements, making it more and more

essential to organize file systems to maximize sequential transfers to and from disk.

Workloads are also evolving rapidly, which changes demands on file systems. In servers, the rising popularity of virtual machines and cloud computing pressure operating systems designers to provide better ways to share storage devices with fair and predictable performance despite variable and mixed workloads. At clients, the increasing popularity of apps and specialized compute appliances are providing new ways for organizing storage: rather than having users organize files into directories, apps and appliances often manage their own storage, providing users with a perhaps very different way of identifying stored objects. For example, rather than requiring users to create different directories for different, related sets of photos into different directories, many photo organizing applications provide an interface that groups related photos into events that may or may not reflect where in the file system the events are stored. Perhaps our reliance on directories for naming and locality will need to be rethought in the coming years.

Exercises

1. Why do many file systems have separate system calls for removing a regular file (e.g., unlink) and removing a directory (e.g., rmdir)?
2. In Figure 13.4, suppose we create a new file z.txt and then unlink work, removing that entry. Draw a figure similar to Figure 13.4 that shows the new contents of the directory.
3. What effect will doubling the block size in the UNIX Fast File System have on the maximum file size?
4. Is there a limit on the maximum size of a file in an extent-based file system? Why or why not?
5. Suppose a variation of FFS includes in each inode 12 direct, 1 indirect, 1 double indirect, 2 triple indirect, and 1 quadruple indirect pointers. Assuming 6 KB blocks and 6-byte pointers
   a. What is the largest file that can be accessed via direct pointers only?
   b. To within 1%, what is the maximum file size this index structure can support?
6. On a Unix or Linux system, use the ls -l command to examine various directories. After the first ten characters that define each file's access permissions, there is a field that indicates the number of hard links to the file. For example, here we have two files, bar with two links and foo with just one.

   drwxr-sr-x 2 dahlin prof 4096 2012-02-03 08:37 bar/
   -rw-r—r— 1 dahlin prof 0 2012-02-03 08:36 foo

   For directories, what is the smallest number of links you can observe? Why?

   For directories, even though regular users cannot make hard links to directories, you may observe some directories with high link counts. Why?
7. In NTFS, a master file table entry maximizes the number of extent pointers it can store by storing extent pointers as a sequence of variable-length records: the first four bits encode the size used to store pointer to the start of the extent and the next four bits encode the size used to store the extent length. To further reduce record size, the extent-start pointer is stored as an offset from the previous extent's pointer. Thus, if we have a 10 block extent starting at block 0x20000 and then a 5 block extent starting at block 0x20050, then the first (absolute) starting address (0x200000) will be stored in three bytes while the second (relative) starting address will be stored in one byte (0x20050 - 0x20000 = 0x50).

   An apparent disadvantage of this approach is that seeking to a random offset in a file requires sequentially scanning all of the extent pointers. Given your understanding of NTFS and disk technology trends, explain why this apparent disadvantage may not be a problem in practice.
8. When user tries to write a file, the file system needs to detect if that file is a directory so that it can restrict writes to maintain the directory's internal consistency.

   Given a file's name, how would you design each file system listed below to keep track of whether each file is a regular file or a directory?
   a. The FAT file system
   b. FFS
   c. NTFS
9. Why would it be difficult to add hard links to the FAT file system?
10. For the FFS file system illustrated in Figure 13.25, what reads and writes of inodes and blocks would occur to create a new file /foo/sparse and write blocks 1 and 2,000,000 of that file. Assume that inodes have pointers for 11 direct blocks, 1 indirect block, 1 double-indirect block, and 1 triple indirect block, and assume 4KB blocks with 4-byte block pointers.
11. Give a formula for the minimum and maximum number of disk blocks that must be read in the UNIX Fast File System to fetch the first block of a file, as a function of the number of "/" characters in the file name (in other words, the depth of the file in the directory tree). Assume that nothing is in the file cache.
12. A web client and web server are running on the same uniprocessor computer. They have an open connection and are ready to send/receive web requests. List a possible sequence of user-mode/kernel-mode boundary crossings (counting one for each direction, and including interrupts) needed for the client to issue a simple web request, the server to receive the request and fetch the data from the file system, and for the server to send the data to the client.

    Assume that there is a low priority background task running on the processor, the current directory is cached, but the requested file is not in the server cache or the file system cache. Also assume that both the request and the requested file data are small (e.g., they fit inside a single disk block). You may assume any of the file systems described in this chapter, provided you label which one you are assuming.

14. Reliable Storage

A stitch in time saves nine. —*English Proverb*

Highly reliable storage is vitally important across a wide range of applications from businesses that need to know that that their billing records are safe to families that have photo albums they would like to last for generations.

So far, we have treated disk and flash as ideal non-volatile storage: stored data will remain forever or until it is overwritten. Physical devices cannot achieve such perfection — they may be defective, they may wear out, or they may be damaged so they may lose some or all of their data.

Unfortunately, the limits of physical devices are not merely abstract concerns. For example, some large organizations have observed annual disk failure rates of 2% to 4%, meaning that an organization with 10,000 disks might expect to see hundreds of failures per year and that important data stored on a single disk by a naive storage system might have more than a 30% chance of disappearing within a decade.

The central question of this chapter is: How can we make a storage system more reliable than the physical devices out of which it is built?

A system is *reliable* if it performs its intended function. Reliability is related to, but different from, availability. A system is *available* if it currently can respond to a request.

In the case of storage, a storage system is reliable as long as it continues to store a given piece of data and as long as its components are capable of reading or overwriting that data. We define a storage system's *reliability* as the probability that it will continue to be reliable for some specified period of time. A storage system is available at some moment if a read or write operation could be completed at that time, and we define a storage system's *availability* as the probability that the system will be available at any given time.

Figure 14.1: The Voyager "Golden Record," a highly reliable but highly unavailable storage device. Photo Credit: NASA.

To see the difference between reliability and availability, consider the highly reliable but highly unavailable storage device shown in Figure 14.1. In the 1970's, the two Voyager spacecraft sent out of our solar system each included a golden record on which various greetings, diagrams, pictures, natural sounds, and music were encoded, as stated on each record by President Carter, as "a present from a small, distant world, a token of our sounds, our science, our images, our music, our thoughts and our feelings." To protect against erosion, the record is encased in an aluminum and uranium cover. This storage device is highly reliable — it is expected to last for many tens of thousands of years in interstellar space — but it is not highly available (at least, not to us).

To take a more pedestrian example, suppose a storage system required each data block to be written to a disk on each of 100 different machines physically distributed across 100 different machine rooms spread across the world. Such a system would be highly reliable, since it would take a spectacular catastrophe to wipe out all of the copies of any data that is stored. It would be highly available for reads, since there are 100 different locations to read from. But it would not be highly available for writes, since new writes cannot complete if any one of 100 machines is unavailable.

Two problems. Broadly speaking, storage systems must deal with two threats to reliability.

- **Operation interruption.** A crash or power failure in the middle of a series of related updates may leave the stored data in an inconsistent state.

  For example, suppose that a user has asked an operating system to move a file from one directory to another:
  > mv drafts/really-important.doc final/really-important.doc

  As we discussed in Chapter 13, such a move may entail many low-level operations: writing the drafts directory file to remove really-important.doc, updating the last-modified time of the drafts directory, growing the final directory's file to include another block of storage to accommodate a new directory entry for really-important.doc, writing the new directory entry to the directory file, updating the file system's free space bitmap to note that the newly allocated block is now in use, and updating the size and last-modified time of the final directory.

  Suppose that the system's power fails when the updates to the drafts directory are stored in non-volatile storage but when the updates to the final directory are not; in that case, the file really-important.doc may be lost. Or, suppose that the operating system crashes after updating the drafts and final directories but before updating the file system's free space bitmap; in that case, the file system will still regard the new block in the final directory as free, and it may allocate that block to be part of some other file. The storage device then ends up with a block that belongs to two files, and updates intended for one file may corrupt the contents of the other file.

- **Loss of stored data.** Failures of non-volatile storage media can cause previously stored data to disappear or be corrupted. Such failures can affect individual blocks, entire storage devices, or even groups of storage devices.

  For example, a disk sector may be lost if it is scratched by a particle contaminating the drive enclosure; a flash memory cell might lose its contents when large numbers of reads of nearby cells disturb its charge; a disk drive can fail completely because bearing wear causes the platters to vibrate too much to be successfully read or

written; or a set of drives might be lost when a fire in a data center destroys a rack of storage servers.

Two solutions. Fortunately, system designers have developed two sets of powerful solutions to these problems, and the rest of the chapter discusses them.

- **Transactions for atomic updates.** When a system needs to make several related updates to non-volatile storage, it may want to ensure that the state is modified atomically: even if a crash occurs the state reflects either all of the updates or none of them. Transactions are a fundamental technique to provide atomic updates of non-volatile storage

  Transactions are simple to implement and to use, and they often have as good or better performance than ad-hoc approaches. The vast majority of widely used file systems developed over the past two decades have used transactions internally, and many applications implement transactions of their own to keep their persistent state consistent.

- **Redundancy for media failures.** To cope with data loss and corruption, storage systems use several forms of redundancy such as checksums to detect corrupted storage and replicated storage to recover from lost or corrupted sectors or disks.

  Implementing sufficient redundancy at acceptably low cost can be complex. For example, a widely used, simple model of RAID (Redundant Array of Inexpensive Disks) paints an optimistic picture of reliability that can be off by orders of magnitude. Modern storage systems often make use of multiple levels of checksums (e.g., both in storage device hardware and file system software), include sufficient redundancy to survive two or more hardware failures (e.g., keeping three copies of a file or two parity disks with a RAID), and rely on software that to detect failures soon after they occur and to repair failures quickly (e.g., background processes that regularly attempt to read all stored data and algorithms that parallelize recovery when a device fails). Systems that fail to properly use these techniques may be significantly less reliable than expected.

14.1 Transactions: Atomic Updates

When a system makes several updates to non-volatile storage and a crash occurs, some of those updates may be stored and survive the crash and others may not. Because a crash may occur without warning, storage systems and applications need to be constructed so that no matter when the crash occurs, the system's non-volatile storage is left in some sensible state.

This problem occurs in many contexts. For example, if a crash occurs while you are installing an update for a suite of applications, upon recovery you would like to be able to use either the old version or the new version, not be confronted with a mishmash of incompatible programs. For example, if you are moving a subdirectory from one location to another when a crash occurs, when you recover you want to see the data in one location or the other; if the subdirectory disappears because of an untimely crash, you will be (justifiably) upset with the operating system designer. Finally, if a bank is moving $100 from Alice's account to Bob's account when a crash occurs, it wants to be certain that upon recovery either the funds are in Alice's account and records show that the transfer is still to be done or that the funds are in Bob's account and the records show that the transfer has occurred.

This problem is quite similar to the critical section problem in concurrency. In both cases, we have several updates to make and we want to avoid having anyone observe the state in an intermediate, inconsistent state. In addition, we have no control when other threads might try to access the state in the first case or when a crash might occur in the second — we must develop a structured solution that works for any possible execution. The solution is similar, too; we want to make the set of updates atomic. However, because we are dealing with non-volatile storage rather than main memory, the techniques for achieving atomicity differ in significant ways.

Transactions extend the concept of atomic updates from memory to stable storage, allowing systems to atomically update multiple persistent data structures.

14.1.1 Ad Hoc Approaches

Until the mid-1990's, many file systems used ad hoc approaches to solving the problem of consistently updating multiple on-disk data structures.

For example, the Unix fast file system (FFS) would carefully control the order that its updates were sent to disk so that if a crash occurred in the middle of a group of updates, a scan of the disk during recovery could identify and repair inconsistent data structures. When creating a new file, for example, FFS would first update the free-inode bitmap to indicate that the previously free inode was now in use. After making sure this update was on disk, it would initialize the new file's inode, clear all of the direct, indirect, double-indirect, and other pointers, set the file length to 0, and set the file's ownership and access control list. Finally, once the inode update was safely on disk, the file system would update the directory to contain an entry for the newly created file, mapping the file's name to its inode.

If a system running FFS crashed, then when it rebooted it would use a program called fsck (file system check) to scan all of the file system's metadata (e.g., all inodes, all directories, and all free space bitmaps) to make sure that all metadata items were consistent. If, for example, fsck discovered an inode that was marked as *allocated* in the free-inode bitmap but that did not appear in any directory entry, it could infer that the inode was part of a file in the process of being created (or deleted) when the crash occurred. Since the create had not finished or the delete had started, fsck could mark the inode as free, undoing the partially completed create (or completing the partially completed delete).

Similar logic was used for other file system operations.

This approach of careful ordering of operations with scanning and repair of on-disk data structures was widespread until the 1990's, when it was largely abandoned. In particular, this approach has three significant problems:

1. **Complex reasoning.** Similar to trying to solve the multi-threaded synchronization problem with just atomic loads and stores, this approach requires reasoning carefully about all possible operations and all possible failure scenarios to make sure that it is always possible to recover the system to a consistent state.
2. **Slow updates.** To ensure that updates are stored in an order that allowed the system's state to be analyzed, file systems are forced to insert sync operations or barriers between dependent operations, reducing the amount of pipelining or parallelism in the stream of requests to storage devices.

   For example, in the file creation example above to ensure that the individual updates hit disk in the required order, the system might suffer three full rotations of the disk to update three on-disk data structures even though those data structures may be quite near each other.
3. **Extremely slow recovery.** When a machine reboots after a crash, it has to scan all of its disks for inconsistent metadata structures.

   In the 1970's and 1980s, it was possible to scan the data structures on most servers' disks in a few seconds or minutes. However, by the 1990's this scanning could take tens of minutes to a few hours for large servers with many disks, and technology trends indicated that scan times would rapidly grow worse.

Although the first two were significant disadvantages of the approach, it was the third that finally made depending on careful ordering and fsck untenable for most file systems. New file systems created since the late 1980's almost invariably use other techniques — primarily various forms of transactions that we discuss in the rest of this section.

---

fsck lives

Although few file systems today rely on scanning disks when recovering from a crash, fsck and other similar programs are often still used as an "emergency fix" when on-disk data structures are corrupted for other reasons (e.g., due to software bug or storage device failure).

Application-level approaches. Although modern file systems often use transactions internally, some standard file system APIs such as the POSIX API provide only weaker abstractions, forcing applications to take their own measures if they want to atomically apply a set of updates. Many use application-level transactions, but some continue to use ad hoc approaches.

For example, suppose that a user has edited several parts of a text file and then wants to save the updated document. The edits may have inserted text at various points in the document, removed text at others, and shifted the remaining text forwards or backwards — even a small insertion or deletion early in the document could ripple through the rest of the file.

If the text editor application were simply to use the updated file in its memory to overwrite the existing file, an untimely crash could leave the file in an incomprehensible state. The operating system and disk schedulers may choose any order to send the updated blocks to non-volatile storage, so after the crash the file may be an arbitrary mix of old and new blocks, sometimes repeating sections of text, sometimes omitting them entirely.

To avoid this problem, the text editor may take advantage of the semantics of the POSIX rename operation, which renames the file called sourceName to be called targetName instead. POSIX promises that if a file named targetName already exists, rename's shift from having targetName refer to the old file to having it refer to the new one will be atomic. (This atomicity guarantee may be provided by transactions within the file system or by ad hoc means.)

Therefore, to update an existing file design.txt, the text editor first writes the updates to a new, temporary file such as #design.txt#. Then it renames the temporary file to atomically replace the previously stored file.

14.1.2 The Transaction Abstraction

Transactions provide a way to atomically update multiple pieces of persistent state.

For example, suppose you are updating a web site and you want to replace the current collection of documents in /server/live with a new collection of documents you have created in /development/ready. You don't want users to see intermediate steps when some of the documents have been updated and others have not — they might encounter broken links or encounter new descriptions referencing old pages or vice versa. Transactional file systems like Windows Vista's TxF (Transactional NTFS) provide an API that lets applications apply all of these updates atomically, allowing the programmer to write something like the pseudo-code in Figure 14.2.

```
ResultCode publish() {
    transactionID = beginTransaction();
    foreach file f in /development/ready that is not in /server/live {
        error = move f from /development/ready to /server/live;
        if (error) {
            rollbackTransaction(transactionID);
            return ROLLED_BACK;
        }
    }

    foreach file f in /server/live that is not in /development/ready {
        error = delete f;
        if (error) {
            rollbackTransaction(transactionID);
            return ROLLED_BACK;
        }
    }

    foreach file f in /development/ready that differs from /server/live {
        error = move f from /development/ready to /server/live;
        if (error) {
            rollbackTransaction(transactionID);
            return ROLLED_BACK;
        }
    }
```

```
    commitTransaction(transactionID);
    return COMMITTED;
}
```

Figure 14.2: Pseudo-code for using a transactional file system.

Notice that a transaction can finish in one of two ways: it can *commit*, meaning all of its updates occur, or it can *roll back* meaning that none of its updates occur.

Here, if the transaction commits, we are guaranteed that all of the updates will be seen by all subsequent reads, but if it encounters and error and rolls back or crashes without committing or rolling back, no reads outside of the transaction will see any of the updates.

More precisely, a *transaction* is a way to perform a set of updates while providing the following *ACID properties*:

- **Atomicity.** Updates are "all or nothing." If the transaction *commits*, all updates in the transaction take effect. If the transaction *rolls back*, then none of the updates in the transaction have any effect.

  In the website update example above, doing the updates within a transaction guarantees that each of the update is only stored or readable if all of the updates are stored and readable.

- **Consistency.** The transaction moves the system from one legal state to another. A system's invariants on its state can be assumed to hold at the start of a transaction and must hold when the transaction commits.

  In the example above, by using a transaction we can maintain the invariant that every link from one document to another on the server references a valid file.

- **Isolation.** Each transaction appears to execute on its own, and is not affected by other in-progress transactions. Even if multiple transactions execute concurrently, for each pair of transactions T and T′, it either appears that T executed entirely before T′ or vice versa.

  By executing the web site update in a transaction, we guarantee that each transaction to read from the web site occurs against either the old set of web pages or the new set, not some mix of the two.

  Of course, if each individual read of an object is in its own transaction, then a series of reads to assemble a web page and its included elements could see the old web page and a mix of old and new elements. If web protocols were changed to allow a browser to fetch a page and its elements in a single transaction, then we could guarantee that the user would see either the old page and elements or the new ones.

- **Durability.** A committed transaction's changes to state must survive crashes. Once a transaction is committed, the only way to change the state it produces is with another transaction.

  In our web update example, the system must not return from the commitTransaction() call until all of the transaction's updates have been safely stored in persistent storage.

Transactions vs. Critical Sections. The ACID properties are closely related to the properties of critical sections. Critical sections provide a way to update state that is atomic, consistent, and isolated but not durable. Adding the durability requirement significantly changes how we implement atomic updates.

Battling terminology

In operating systems, we use the term *consistency* in two ways. In the context of critical sections and transactions, we use "consistency" to refer to the idea of a system's invariants being maintained (e.g., "are my data structures consistent?") In the context of distributed memory machines and distributed systems, we use "consistency" to refer to the memory model — the order in which updates can become visible to reads (e.g., "are my system's reads at different caches sequentially consistent?").

Where there is potential confusion, we will use the terms *transaction consistency* or *memory model consistency*.

14.1.3 Implementing Transactions

The challenge with implementing transaction is that we want a group of related writes to be atomic, but for persistent storage hardware like disks and flash, the atomic operation is a single-sector or single-page write. Therefore, we must devise a way for a group of related writes to take effect when a single-sector write occurs.

If a system simply starts updating data structures in place, then it is vulnerable to a crash in the middle of a set of updates: the system has neither the complete set of old items (to roll back) nor a complete set of new items (to commit), so an untimely crash can force the system to violate atomicity.

Instead, a transactional system can persistently store all of a transaction's *intentions*, the updates that will be made if the transaction commits, in some separate location of persistent storage. Only when all intentions are stored and the transaction commits should the system begin overwriting the target data structures; if the overwrites are interrupted in the middle, then on recovery the system can complete the transaction's updates using the persistently stored intentions.

Redo Logging

A common and very general way to implement transactions is redo logging. *Redo logging* uses a persistent log for recording intentions and executes a transaction in four stages:

1. **Prepare.** Append all planned updates to the log.

   This step can happen all at once, when the transaction begins to commit, or it can happen over time, appending new updates to the log as the transaction executes. What is essential is that all updates are safely stored in the log before proceeding to the next step.

2. **Commit.** Append a commit record to the log, indicating that the transaction has committed.

   Of course, a transaction may roll back rather than commit. In this case, a roll-back record may be placed in the log to indicate that the transaction was abandoned. Writing a roll-back record is optional, however, because a transaction will only be regarded as committed if a commit record appears in the log.

3. **Write-back.** Once the commit record is persistent in the log, all of a transaction's updates may be written to their target locations, replacing old values with new ones.

4. **Garbage collect.** Once a transaction's write-back completes, its records in the log may be garbage collected.

The moment in step 2 when the sector containing the commit record is successfully stored is the *atomic commit*: before that moment, the transaction may safely be rolled back; after that moment, the transaction must take effect.

Recovery. If a system crashes in the middle of a transaction, it must execute a recovery routine before processing new requests. For redo logging, the recovery routine is simple: scan sequentially through the log, taking the following actions for each type of record:

1. **Update record for a transaction.** Add this record to a list of updates planned for the specified transaction.
2. **Commit record for a transaction.** Write-back all of the transaction's logged updates to their target locations.
3. **Roll-back record for a transaction.** Discard the list of updates planned for the specified transaction.

When the end of the log is reached, the recovery process discards any update records for transactions that do not have commit records in the log.

Example. Consider, for example, a transaction that transfers $100 from Tom's account to Mike's account. Initially, as Figure 14.3-(a) shows, data stored on disk and in the volatile memory cache indicates that Tom's account has $200 and Mike's account has $100.

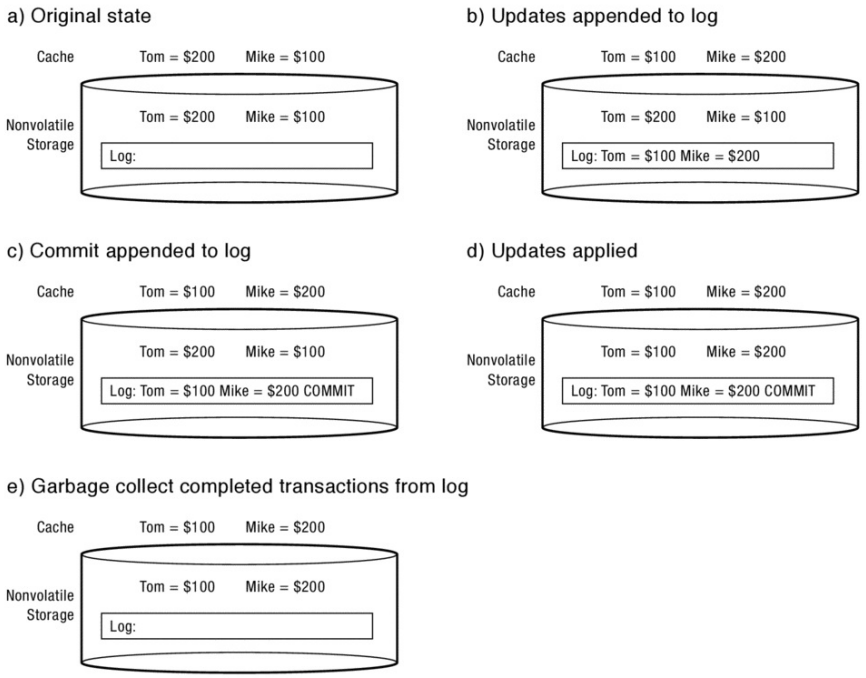


Figure 14.3: Example transaction with redo logging.

Then, the cached values are updated and the updates are appended to the non-volatile log (b). At this point, if the system were to crash, the updates in cache would be lost, the updates for the uncommitted transaction in the log would be discarded, and the system would return to its original state.

Once the updates are safely in the log, the commit record is appended to the log (c). This commit record should be written atomically based on the properties of the underlying hardware (e.g., by making sure it fits on a single disk sector and putting a strong checksum on it). This step is the atomic commit: prior to the successful storage of the commit record, a crash would cause the transaction to roll back; the instant the commit record is persistently stored, the transaction has committed and is guaranteed to be visible to all reads in the future. Even if a crash occurs, the recovery process will see the committed transaction in the log and apply the updates.

Now, the records in persistent storage for Tom and Mike's accounts can be updated (d).

Finally, once Tom and Mike's accounts are updated, the transaction's records in the log may be garbage collected (e).

Implementation details. A few specific techniques and observations are important for

providing good performance and reliability for transactions with redo logs.

- **Logging concurrent transactions.** Although the previous example shows a single transaction, multiple transactions may be executing at once. In these cases, each record in the log must identify the transaction to which it belongs.
- **Asynchronous write-back.** Step 3 of a transaction (*write-back*) can be asynchronous — once the updates and commit are in the log, the write-back can be delayed until it is convenient or efficient to perform it.

  This flexibility yields two advantages. First, it minimizes the latency from when a transaction calls commit to when the call returns. As soon as the commit is appended to the sequential log, the call can return. Second, the throughput for write-back is higher because the disk scheduler can operate on large batches of updates.

  Two things limit the maximum write-back delay, but both are relatively loose constraints. First, larger write-back delays mean that crash recovery may take longer because there may be more updates to read and apply from the log. Second, the log takes space in persistent storage, which may in some cases be constrained.
- **Repeated write-backs are OK.** Some of the updates written back during recovery may already have been written back before the crash occurred. For example, in Figure 14.4 all of the records from the persistent log-head pointer to the volatile one have already been written back, and some of the records between the volatile log-head pointer may have been written back.

  It is OK to reapply an update from a redo log multiple times because these updates are (and must be) *idempotent* — they have the same effect whether executed once or multiple times. For example, if a log record says "write 42 to each byte of sector 74" then it doesn't matter whether that value is written once, twice, or a hundred times to sector 74.

  Conversely, redo log systems cannot permit non-idempotent records such as "add 42 to each byte in sector 74."
- **Restarting recovery is OK.** What happens if another crash occurs during recovery? When the system restarts, it simply begins recovery again. The same sequence of updates to committed transactions will be discovered in the log, and the same write-backs will be issued. Some of the write-backs may already have finished before the first crash or during some previous, but repeating them causes no problems.
- **Garbage collection constraints.** Once write-back completes and is persistently stored for a committed transaction, its space in the log can be reclaimed.

  For concreteness, Figure 14.4 illustrates a transaction log with an area of the log that is in use, an area that is no longer needed because it contains only records for transactions whose write-backs have completed, and an area that is free.

  In this example, the system's volatile memory maintains pointers to the head and tail of the log; new transaction records are appended to the log's tail and cached in volatile memory; a write-back process asynchronously writes pending write-backs for committed transactions to their final locations in persistent storage; and a garbage collection process periodically advances a persistent log-head pointer so that recovery can skip at least some of the transactions whose write-backs are complete.
- **Ordering is essential.** It is vital to make sure that all of a transaction's updates are on disk in the log before the commit is, that the commit is on disk before any of the write-backs are, and that all of the write-backs are on disk before a transaction's log records are garbage collected.

  In Linux, an application can call sync() or fsync() to tell the operating system to force buffered writes to disk. These calls return only once the updated blocks are safely stored. Within the operating system, a request can be tagged with a BIO_RW_BARRIER tag, which tells the device driver and storage hardware to ensure that all preceding writes and no subsequent ones are stored before the tagged request is.

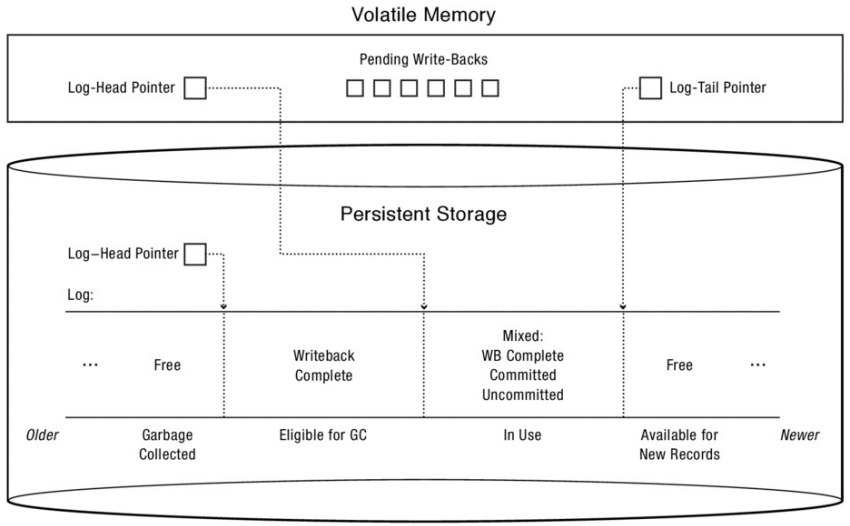


Figure 14.4: Volatile and persistent data structures for a transactional system based on a replay log.

EXAMPLE: New writes vs. garbage collection. Suppose we have a circular log organized like the one in Figure 14.4. We must ensure that new records do not overwrite records that we may read during recovery, so we must ensure that the log-tail does not catch the log-head. But there are two log-heads, one in volatile memory and another in persistent storage. Which log-head represents the barrier that the log-tail must not cross?

ANSWER: The log-tail must not catch the persistent log-head pointer. Even though the records between the persistent and volatile log-heads have already been written back, during crash recovery, the recovery process will begin reading the log from the location indicated by the persistent log-head pointer. As long as the records are intact, recovering

from the persistent log-head pointer rather than the volatile one may waste some work, but it will not affect correctness. □

Undo logging

Although transactions are often implemented with redo logging in which updates and the commit are written to the log and then the updates are copied to their final locations, transactions can also be implemented with *undo logging*.

To update an object, a transaction first writes the *old* version of the object to the log. It then writes the new version to its final storage location. When the transaction completes, it simply appends *commit* to the log. Conversely, if the transaction rolls back, the updates are undone by writing the old object versions to their storage locations.

The recovery process takes no action for committed transactions it finds in the log, but it undoes uncommitted transactions by rewriting the original object versions stored in the log.

Undo logging allows writes to objects to be sent to their final storage locations when they are generated and requires them to be persistently stored before a transaction is committed. This pattern is similar to update-in-place approaches, so in some cases it may be easier to add undo logging than redo logging to legacy systems. On the other hand, for storage systems like disks whose sequential bandwidth dominates their random I/O performance, undo logging may require more random I/Os before a transaction is committed (hurting latency) and by writing each transaction's updates immediately, it gives up chances to improve disk-head scheduling by writing large numbers of transaction updates as a batch.

Undo/redo logging stores both the old and new versions of an object in the log. This combination allows updated objects to be written to their final storage locations whenever convenient, whether before or after the transaction is committed. If the transaction rolls back, any modified objects can be restored to the proper state, and if the system crashes, any committed transactions can have their updates redone and any uncommitted transactions can have their updates undone.

Isolation and concurrency revisited. Redo logging provides a mechanism for atomically making multiple updates durable, but if there are concurrent transactions operating on shared state, we must also ensure isolation — each transaction must appear to execute alone, without interference from other transactions.

A common way to enforce isolation among transactions is *two-phase locking*, which divides a transaction into two phases. During the *expanding* phase, locks may be acquired but not released. Then, in the *contracting* phase, locks may be released but not acquired. In the case of transactions, because we want isolation and durability, the second phase must wait until after the transaction commits or rolls back so that no other transaction sees updates that later disappear.

As we discussed in Chapter 6, two phase locking ensures a strong form of isolation called serializability. *Serializability* across transactions ensures that the result of any execution of the program is equivalent to an execution in which transactions are processed one at a time in some sequential order. So, even if multiple transactions are executed concurrently, they can only produce results that they could have produced had they been executed one at a time in some order.

Although acquiring multiple locks in arbitrary orders normally risks deadlock, transactions provide a simple solution. If a set of transactions deadlocks, one or more of the transactions can be forced to roll back, release their locks, and restart at some later time.

Multiversion concurrency control

An alternative to enforcing transaction isolation with locks is to enforce it with multiversion concurrency control. In *multiversion concurrency control* each write of an object x creates new version of x, the system keeps multiple versions of x and directs each read to a specific version of x. By keeping multiple versions of objects, the system can allow transaction A to read a version of x that has been overwritten by transaction B even if B needs to be serialized after A.

Several multiversion concurrency control algorithms ensure serializability. A simple one is multiversion timestamp ordering (MVTO). MVTO processes concurrent transactions, enforces serializability, never blocks a transaction's reads or writes, but may cause a transaction to roll back if it detects that a read of a later transaction (based on the serializable schedule MVTO is enforcing) was executed before — and therefore did not observe — the write of an earlier transaction (in serialization order).

MVTO assigns each transaction T a logical timestamp. Then, when T writes an object x, MVTO creates a new version of x labeled with T 's timestamp t_T, and when T reads an object y, MVTO returns the version of y, y_V with the highest timestamp that is at most T 's timestamp; MVTO also makes note that y_V was read by transaction t_T. Finally, when T attempts to commit, MVTO blocks the commit until all transactions with smaller timestamps have committed or aborted.

MVTO rolls back a transaction rather than allowing it to commit in three situations. First, if MVTO aborts any transaction, it removes the object versions written by that transaction and rolls back any transactions that read those versions. Notice that a transaction that reads a version must have a higher timestamp than the one that wrote it, so no committed transactions need to be rolled back.

Second, if a transaction T writes an object that has already been read by a later transaction T′ that observed the version immediately prior to T 's write, T MVTO rolls back T . It does this because if T were to commit, T′'s read must return T 's write, but that did not occur.

Third, if MVTO garbage collects old versions and transaction T reads an object for which the last write by an earlier transaction has been garbage collected, then MVTO rolls back T .

Relaxing isolation

In this book we focus on the strong and relatively simple isolation requirement of serializability: no matter how much concurrency there is, the system must ensure that the results of any execution of the program is equivalent to an execution in which transactions are processed one at a time in some sequence. However, strong isolation requirements sometimes force transactions to block (e.g., when waiting to acquire locks) or roll back (e.g., when fixing a deadlock or encountering a "late write" under multiversion concurrency control).

Relaxing the isolation requirement can allow effectively higher levels of concurrency by reducing the number of cases in which transactions must block or roll back. The cost, of course, is potentially increased complexity in reasoning about concurrent programs, but several relaxed isolation semantics have proven to be sufficiently strong to be widely used.

For example, *snapshot isolation* requires each transaction's reads appear to come from a snapshot of the system's committed data taken when the transaction starts. Each transaction is buffered until the transaction commits, at which point the system checks all of the transaction's updates for *write-write conflicts*. A write-write conflict occurs if transaction T reads an object o from a snapshot at time t_{start} and tries to commit at time t_{commit} but some other transaction T′ commits an update to o between T 's read at t_{start} and T 's attempted commit at t_{commit}. If a write-write conflict is detected for any object being committed by T , T is rolled back.

Snapshot isolation is weaker than serializability because each transaction's reads logically happen at one time and its writes logically happen at another time. This split allows, for example, *write skew anomalies* where one transaction reads object x and updates object y and a concurrent transaction reads object y and updates object x. If there is some constraint between x and y, it may now be violated. For example, if x and y represent the number of hours two managers have assigned you to work on each of two tasks with a constraint that $x + y \leq 40$. Manager 1 could read x = 15 and y = 15, attempt to assign 10 more hours of work on task x, and verify that $x + y = 25 + 15 \leq 40$. In the mean time manager 2 could read x = 15 and y = 15, attempt to assign 10 more hours of work on task y, verify that that $x + y = 15 + 25 \leq 40$, and successfully commit the update, setting y = 25. Finally, manager 1 could successfully commit its update, set x = 25, and ruin your weekend.

Performance of redo logging. It might sound like redo logging will impose a significant performance penalty compared to simply updating data in place: redo logging writes each update twice — first to the log and then to its final storage location.

Things are not as bad as they initially seem. Redo logging can have excellent performance — often better than update in place — especially for small writes. Four factors allow efficient implementations of redo logging:

- **Log updates are sequential.** Because log updates are sequential, appending to the log is fast. With spinning disks, large numbers of updates can be written as a sequential stream without seeks or rotational delay once the write begins. Many high-performance systems dedicate a separate disk for logging so that log appends never require seeks. For flash storage, sequential updates are often significantly faster than random updates, though the advantage is not as pronounced.
- **Write-back is asynchronous.** Because write-back can be delayed until some time after a transaction has been committed, transactions using redo logs can have good response time (because the transaction commit only requires appending a commit record to the log) and can have good throughput (because batched write-backs can be scheduled more efficiently than individual or small groups of writes that must occur immediately).
- **Fewer barriers or synchronous writes are required.** Some systems avoid using transactions by carefully ordering updates to data structures so that they can ensure that if a crash occurs, a recovery process will be able to scan, identify, and repair inconsistent data structures. However, these techniques often require large number of barrier or synchronous write operations, which reduce opportunities to pipeline or efficiently schedule updates.

  In contrast, transactions need a relatively small number of barriers: one after the updates are logged and before the commit is logged, another after the commit is logged but before the transaction is reported as successful (and before write-backs begin), and one after a transaction's write-backs complete but before the transaction's log entries are garbage collected.
- **Group commit.** *Group commit* can further improve transaction performance. Group commit combines a set of transaction commits into one log write to amortize the cost of initiating the write (e.g., seek and rotational delays). Group commit techniques can also be used to reduce the number of barrier or sync operations needed to perform a group of transactions.

EXAMPLE: Performance of small-write transactions. Suppose you have a 1 TB disk that rotates once every 10 ms, that has a maximum sustained platter transfer rate of 50 MB/s for inner tracks and 100 MB/s for outer tracks, and that has a 5 ms average seek time, a 0.5 ms minimum seek time, and a 10 ms maximum seek time.

Consider updating 100 randomly selected 512-byte sectors; assume that the updates must be ordered for safety (e.g., update i must be on disk before update i + 1 is applied).

Compare the total time to complete these updates using a simple update in place approach with the cost when using transactions implemented with a redo log.

ANSWER: Using a simple update in place approach, we need to use FIFO scheduling to ensure updates hit the disk in order, so the time for each update is approximately:

$$\begin{aligned} \text{time per update} &= \text{average seek time} + 0.5 \text{ rotation time} + \text{transfer time} \\ &= 5 \text{ ms} + 5 \text{ ms} + \text{transfer time} \end{aligned}$$

Transfer time will be at most 512 / (50×10^6) seconds, which will be negligible compared to the other terms. Thus, we have **10 ms per request** or **1 s for 100 requests** for update in place.

For transactions, we first append the 100 writes to the log. We will conservatively assume that each update consumes 2 sectors (one for the data and the other for metadata indicating the transaction number and the target sector on disk). So, assuming that the disk head is at a random location when the request arrives, our time to log the requests is:

time to write log = average seek time + 0.5 rotation time + transfer time

= 5 ms + 5 ms + transfer time

= 5 ms + 5 ms + (200 × 512) / (100×10^6) s

= 11.0 ms

Next, we need to append the commit record to the transaction. If the disk hardware supports a barrier instruction to enforce ordering of multiple in-progress requests, the operating system can issue this request along with the 100 writes. Here, we will be conservative and assume that the system does not issue the commit's write until after the 100 writes in the body of the transaction are in the log. Thus, we will likely have to wait one full revolution of the disk to finish the commit: *10 ms.*

Finally, we need to write the 100 writes to their target locations on disk. Unlike the case for update in place, ordering does not matter here, so we can schedule them and write them more efficiently. Estimating this time takes engineering judgment, and different people are likely to make different estimates. For this example, we will assume that the disk uses a variant of shortest service time first (SSTF) scheduling in which the scheduler looks at the four requests on the next nearest tracks and picks the one with the shortest predicted seek time + rotational latency from the disk head's current position. Because the scheduler gets to choose from four requests, we will estimate that the average rotational latency will be one fourth of a rotation, 2.5 ms. This may be conservative since it ignores the fact that request i will always remove from the four requests being considered the one that would have been rotationally farthest away if it were an option for request i + 1. Because we initially have 100 requests and because we are considering the four requests on the nearest tracks, the farthest seek should be around 4% of the way across the disk, and the average one to a member of the group being considered should be around 2%. We will estimate that seeking 2-4% of the way across disk costs twice the minimum seek time: 1 ms.

Putting these estimates for write-back time together, the write-backs of the 100 sectors should take about:

per-request write-back time = seek time + rotational latency

= 1.0 ms + 2.5 ms

= 3.5 ms

giving us a total of 350 ms for 100 requests.

Adding the logging, commit, and write-back times, we have:

total write time = log time + commit time + write-back time

= 11.0 ms + 10 ms + 350 ms

= **371 ms**

The transactional approach is almost three times faster even though it writes the data twice and even though it provides the stronger atomic-update semantics. □

EXAMPLE: For the same two approaches, compare the response time latency from when a call issuing these requests is issued until that call can safely return because all of the updates are durable.

ANSWER: The time for **update in place** is the same as above: **1 s**. The time for the **transactional approach** is the time for the first two steps: logging the updates and then logging the commit: *10.24 ms + 10 ms =* ***20.24 ms***. □

Although small writes using redo logging may actually see performance benefits compared to update in place approaches, large writes may see significant penalties.

EXAMPLE: Performance of large-write transactions. Considering the same disk and approaches as in the example above, compare the total time to for 100 writes, but now assume that each of the 100 writes updates a randomly selected 1 MB range of sequential sectors.

ANSWER: For the update in place approach, the time for each update is approximately *average seek time + 0.5 rotation time + transfer time = 5 ms + 5 ms + transfer time.* We will assume that the bandwidth for an average transfer is 75 MB/s — between the 50 MB/s and 100 MB/s inner and out tracks' transfer rates. Thus, we estimate the average transfer time to be 100 MB/75 MB/s = 1.333 s, giving a total time of *.005 s + .005 s +*

1.333 s = 1.343 s per request and *134.3 s* for 100 requests.

For the transactional approach, our time will be *time to log updates + time to commit + time to write back*.

For logging the updates, we assume a reasonably efficient encoding of metadata that makes the size of the metadata for a 100 MB sequential update negligible compared to the data. Thus, logging the data will take *seek time + rotational latency + transfer time = 5 ms + 5 ms + 100 × 100 MB/100 MB/s = .005 s + .005 s + 100 s ≈ 100 s*.

Writing the commit adds another 10 ms as in the above example.

Finally, as above, doing the write-backs *estimated scheduled seek time + estimated scheduled latency + transfer time = 1.0 ms + 2.5 ms + 100 MB/75 MB/s*, giving a total of *1.337 s* per request and **133.7 s** for 100 requests.

Adding the data logging, commit, and write-back times together, the **transactional approach takes about 233 s** while the **update in place approach takes about 134 s**. In this case, transactions do impose a significant cost, nearly doubling the total time to process these updates. □

EXAMPLE: Now compare the latency from when the call making the 100 writes is issued until it may safely return.

ANSWER: Under the update in place approach, we can only return when everything is written, while under the transactional approach, we can return once the commit is complete. Thus, we have comparable times: **134 s for update in place** and **100 s for transactions**. □

One way to reduce transaction overheads for large writes is to add a level of indirection: write the large data objects to a free area of the disk, but not in the circular log. Then, the update in the log just needs to be a reference to that data rather than the data itself. Finally, after the transaction commits, perform the write-back by updating a pointer in the original data structure to point to the new data.

14.1.4 Transactions and File Systems

File systems must maintain internal consistency when updating multiple data structures. For example, when a file system like FFS creates a new file, it may need to update the file's inode, the free inode bitmap, the parent directory, the parent directory's inode, and the free space bitmap. If a crash occurs in the middle of such a group of updates, the file system could be left in an inconsistent state with, say, the new file's inode allocated and initialized but without an entry in the parent directory.

As discussed in Section 14.1.1, some early file systems used ad-hoc solutions such as carefully ordering sequences of writes and scanning the disk to detect and repair inconsistencies when restarting after a crash. However, these approaches suffered from complexity, slow updates, and — as disk capacity grew — unacceptably slow crash recovery.

To address these problems, most modern file systems use transactions.

Traditional file systems. Transactions are added to traditional, update-in-place file systems like FFS and NTFS using either *journaling* or *logging* .

- **Journaling.** Journaling file systems apply updates to the system's metadata via transactions, but they update the contents of users' files in place.

  By protecting metadata updates, these systems ensure consistency of their persistent data structures (e.g., updates to inodes, bitmaps, directories, and indirect blocks). Journaling file systems first write metadata updates to a redo log, then commit them, and finally write them back to their final storage locations.

  Updates to the contents of regular (non-directory) files are not logged, they are applied in place. This avoids writing file updates twice, which can be expensive for large updates. On the other hand, updating file contents in place means that journaling file systems provide few guarantees when a program updates a file: if a crash occurs in the middle of the update, the file may end up in an inconsistent state with some blocks but not others updated. If a program using a journaling file system requires atomic multi-block updates, it needs to provide them itself.
- **Logging.** Logging file systems simply include all updates to disk — both metadata and data — in transactions.

Today, journaling file systems are common: Microsoft's NTFS, Apple's HFS+, and Linux's XFS, JFS, and ReiserFS all use journaling; and Linux's ext3 and ext4 use journaling in their default configurations.

Logging file systems are also widely available, at least for Linux. In particular, Linux's ext3 and ext4 file systems can be configured to use either journaling or logging.

Copy-on-write file systems. Copy-on-write file systems like the open source ZFS are designed from the ground up to be transactional. They do not overwrite data in place; updating the root inode or ZFS uberblock is an atomic action that commits a set of updates.

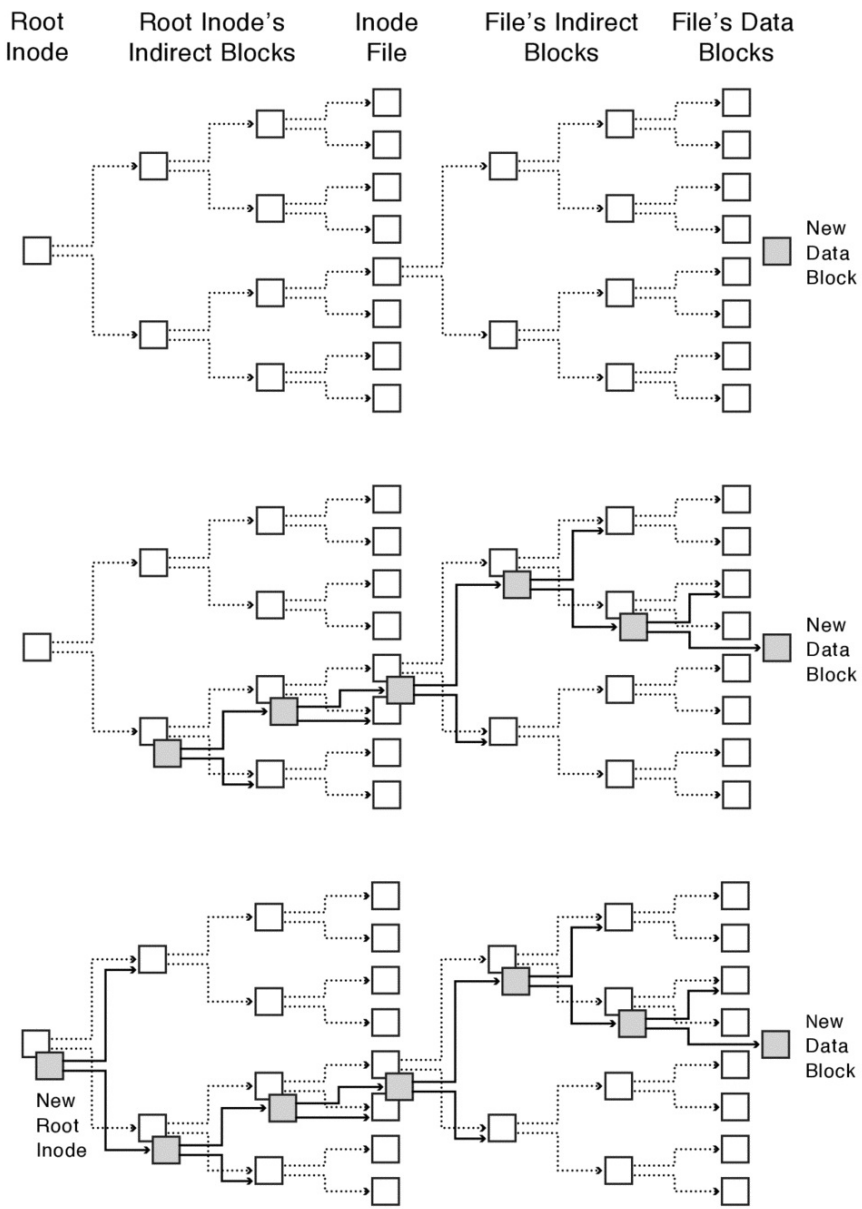


Figure 14.5: In a copy-on-write file system, intermediate states of an update such as (top) and (center) are not observable; they atomically take effect when the root inode is updated (bottom).

For example, suppose we update just a file's data block or just its indirect blocks, its inode, and the indirect blocks for the inode file, leaving the state as shown in Figure 14.5(a) or (b). If the system were to crash in such an intermediate state, before the root inode is updated, none of these changes would be included in the file system's tree, and they would have no effect. Only when the root inode is updated as in Figure 14.5(c) do all of these changes take effect at once.

The implementations of ZFS and other copy-on-write file systems often add two performance optimizations.

- **Batch updates.** Rather than applying each update individually, ZFS buffers several seconds worth of updates before writing them to stable storage as a single atomic group.

  Batching yields two advantages.

  First, it allows the system to transform many small, random writes into a few large, sequential writes, which improves performance for most storage devices including individual magnetic disks, RAIDs (Redundant Arrays of Inexpensive Disks), and even some flash storage devices.

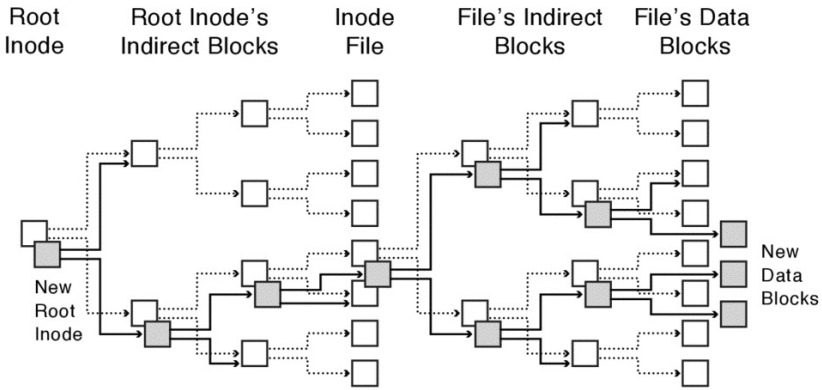


Figure 14.6: With batch updates in a COW file system, updates of inodes and indirect blocks are amortized across updates of multiple data blocks.

Second, not only does batching make writing each block more efficient, it actually reduces how many blocks must be written by coalescing multiple updates of the same indirect blocks and inodes. For example, Figure 14.6 illustrates how updates of inodes and indirect blocks are amortized across updates of multiple data blocks.

- **Intent log.** ZFS typically accumulates several seconds of writes before performing a large batch update, but some applications need immediate assurance that their

updates are safely stored on non-volatile media. For example, when a word processor's user saves a file, the program might call fsync() to tell the file system to make sure the updates are stored on disk. Forcing the user to wait several seconds to save a file is not acceptable.

ZFS's solution is the ZFS Intent Log (ZIL), which is essentially a redo log. The ZIL is a linked list of ZFS blocks that contain updates that have been forced to disk but whose batch update may not yet have been stored. The ZIL is replayed when the file system is mounted.

ZFS includes several optimizations in the ZIL implementation. First, by default writes are buffered and committed in their batch update without being written to the ZIL; only writes that are explicitly forced to disk are written to the ZIL. Second, the ZIL may reside on a separate, dedicated logging device; this allows us to use a fast device (e.g., flash) for the ZIL and slower, high-capacity devices (e.g., disks) for the main pool; if no separate ZIL device is provided, ZFS uses the main block pool for the ZIL. Finally, the contents of small data writes are included in the ZIL's blocks directly, but the contents of larger writes are written to separate blocks that are referenced by the ZIL; then, the subsequent batch commit can avoid rewriting those large blocks by updating metadata to point to the copies already on disk.

14.2 Error Detection and Correction

Because data storage hardware is imperfect, storage systems must be designed to detect and correct errors. Storage systems take a layered approach:

- Storage hardware detects many failures with checksums and device-level checks, and it corrects small corruptions with error correcting codes
- Storage systems include redundancy using RAID architectures to reconstruct data lost by individual devices
- Many recent file systems include additional end-to-end correctness checks

These techniques are essential. Essentially all persistent storage devices include internal redundancy to achieve high storage densities with acceptable error rates. This internal redundancy is insufficient on its own. Storage systems for important data add additional redundancy for error correction, and it is hard to think of a significant file system developed in the last decade that does not include higher-level checksums.

Though essential and widespread, there are significant pitfalls in designing and using these techniques. In our discussions, we will point out issues that, if not handled carefully, can drastically reduce reliability.

The rest of this section examines error detection and correction for persistent storage, starting with the individual storage devices, then examining how RAID replication helps tolerate failures by individual storage devices, and finally looking at the end-to-end error detection in many recent file systems.

14.2.1 Storage Device Failures and Mitigation

Storage hardware pushes the limits of physics, material sciences, and manufacturing processes to maximize storage capacity and performance. These aggressive designs leave little margin for error, so manufacturing defects, contamination, or wear can cause stored bits to be lost.

Individual spinning disks and flash storage devices exhibit two types of failure. First, isolated disk sectors or flash pages can lose existing data or degrade to the point where they cannot store new data. Second, an entire device can fail, preventing access to all of its sectors or pages. We discuss each of these in turn to understand the problems higher-level techniques need to deal with.

Sector and Page Failures

Disk *sector failures* occur when data on one or more individual sectors of a disk are lost, but the rest of the disk continues to operate correctly. *Flash page failures* are the equivalent for flash pages.

Storage devices use two techniques to mitigate sector or page failures: error correcting codes and remapping.

What causes sector or page failures?

For spinning disks, permanent sector failures can be caused by a range of faults such as pits in the magnetic coating where a contaminant flaked off the surface, scratches in the coating where a contaminant was dragged across the surface by the head, or smears of machine oil across some sectors of a disk surface.

Transient sector faults occur when a sector's stored data is corrupted but new data can still be successfully written to and read from the sector. These can be caused by factors such as write interference where writes to one track disturb bits stored on nearby tracks and "high fly writes" where the disk head gets too far from the surface, producing magnetic fields too weak to be accurately read.

For flash storage, permanent page failures can be caused by manufacturing defects or by wear-out when a page experience a large number of write/erase cycles.

Transient flash storage failures can be caused by: (i) write disturb errors where charging one bit also causes a nearby bit to be charged; (ii) read disturb errors where repeatedly reading one page changes values stored on a nearby page; (iii) over-programming errors where too high a voltage is used to write a cell, which may cause incorrect reads or writes; and (iv) data retention errors where charge may leak out of or into a flash cell over time, changing its value. Wear-out from repeated write/erase cycles can make devices more susceptible to data retention errors.

Mitigation: Error correcting codes. *Error correcting codes* deal with failures when some of the bits in a sector or page are corrupted. When the device stores data, it encodes the

data with additional redundancy. Then, if a small number of bits are corrupted in a sector or page being read, the hardware automatically corrects the error, and the read successfully completes. If the damage is more extensive, then with high likelihood the read fails and returns an error code. Being told that the device has lost data is not a perfect solution, but it is better than having the device silently return the wrong data.

Manufacturers balance storage space overheads against error correction capabilities to achieve acceptable advertised sector or page failure rate, typically expressed as the expected number of bits that can be read before encountering an unreadable sector or page. In 2011, advertised disk and flash *non-recoverable read error* rates typically range between one sector/page per 10^{14} to 10^{16} bits read. The non-recoverable read error rate is sometimes called the *bit error rate*.

Mitigation: Remapping. Disks and flash are manufactured with some number of spare sectors or pages so that they can continue to function despite some number of permanent sector or page failures by remapping failed sectors or pages to good ones. Before shipping hardware to users, manufacturers scan devices to remap bad sectors or pages caused by manufacturing defects. Later, if additional permanent failures are detected, the operating system or device firmware can remap the failed sectors or pages to good ones.

Pitfalls. Although devices' non-recoverable read rate specifications are helpful, designers must avoid a number of common pitfalls:

- **Assuming that non-recoverable read error rates are negligible.** Storage devices' advertised error rates sound impressive, but with the large capacities of today's storage, these error rates are non-negligible. For example, if you completely read a 2 TB disk with a bit error rate of one sector per 10^{14} bits, there may be more than a 10% chance of encountering at least one error.
- **Assuming non-recoverable read error rates are constant.** Although a device may specify a single number as its unrecoverable read error rate, many factors can affect the rate at which such errors manifest. A given device's actual bit error rate may depend on its load (e.g., some faults may be caused by device activity), its age (e.g., some faults may become more likely as a device ages), or even its specific workload (e.g., faults in some sectors or pages may be caused by reads or writes to nearby sectors or pages).
- **Assuming independent failures.** Errors may be correlated in time or space: finding an error in one sector may make it more likely that you will find one in a nearby sector or that you will to find a fault in another sector soon.
- **Assuming uniform error rates.** The relative contributions of different causes of non-recoverable read errors can vary across models and different generations or production runs of the same model. For example, one model of disk drive might have many of its sector read errors caused by contaminants damaging its recording surfaces while another model might have most of its errors caused by write interference where writes to one track perturb data stored on nearby tracks. The first might see its error rate rise over time, while the second might have an error rate that increases as its write/read ratio increases.

  Failure rates can even vary across different individual devices. If you deploy several outwardly identical disks, some may exhibit tens of non-recoverable read errors in a year, while others operate flawlessly.

EXAMPLE: Unrecoverable read errors. Suppose that the nearly-full 500 GB disk on your laptop has just stopped working. Fortunately, you have a recent, full backup on a 500 GB USB drive with an unrecoverable read error rate of one sector per 10^{14} bits read. Estimate the probability of successfully reading the entire USB backup disk when restoring your data to a replacement laptop disk.

ANSWER: We need to read 500 GB, so the expected number of failures is 500 GB × 8 × (10^9 bits/GB) × (10^{-14} errors/bit) = 0.04. The probability of encountering at least one failure might be a bit lower than that (since we may encounter multiple failures as we scan the entire disk), but there appears to be a chance of at least a few percent that the restoration will not be fully successful.

We can approach the problem in a slightly different way by interpreting the unrecoverable read rate as meaning that each bit has a 10^{-14} chance of being wrong and that failures are independent (both somewhat dubious assumptions, but probably OK for a ballpark estimate). Then each bit has a $1 - 10^{-14}$ chance of being correct, and the chance of reading all bits successfully is $P_S = (1 - 10^{-14})^{8\times500\times10^9} = 0.9608$. Under this calculation, we estimate that there is slightly less than a 4% chance of encountering a failure during the full-disk read of the backup disk.

As noted in the sidebar, these calculations ignore some important factors, so the results may not be precise. But, even if they are off by as much as an order of magnitude, then it is still reasonable to conclude that the rate of non-recoverable read errors is likely to be non-negligible. □

Note that the impact of a small number of lost sectors may be modest (e.g., the backup software succeeds in restoring all but a file or two) or it may be severe (e.g., no data is restored). For example, if the sector failure corrupts the root directory, a significant fraction of the data may be lost.

Device Failures

A *full disk failure* occurs when a device stops being able to service reads or writes to all disk sectors; a *full flash drive failure* is the equivalent for a flash device.

What causes whole-device failures?

Disk failures can be caused by a range of faults such as a disk head being damaged, a capacitor failure or power surge that damages the electronics, or mechanical wear-out that makes it difficult for the head to stay centered over a track.

Common causes of flash device failures include wear-out, when enough individual pages fail that the device runs out of spare pages to use for remapping, and failures of the device's electronics such as having a capacitor fail.

When a whole device fails, the host computer's device driver will detect the failure, and reads and writes to the device will return error codes rather than, for example, returning incorrect data. This explicit failure notification is important because it reduces the amount of cross-device redundancy needed to correct failures.

Full device failure rates are typically characterized by an *annual failure rate*, the fraction of disks expected to fail each year, or its inverse, the *mean time to failure (MTTF)*. In 2011, specified annual failure rates (or MTTFs) for spinning disks typically range from 0.5% (1.7×10^6 hours) to 0.9% (1×10^6 hours); specified failure rates for flash solid state drives are similar.

Pitfalls. Storage system designers must consider several pitfalls when considering advertised device failure rates.

- **Relying on advertised failure rates.** Studies across several large collections of spinning disks have found significantly variability in failure rates. In these studies, many systems experienced failure rates of 2%, 4%, or higher, despite advertised failure rates of less than 1%.

  Some of the discrepancy may be due to different definitions of "failure" by manufacturers and users, some may be due to challenging field conditions, and some may be due to the limitations of the accelerated-aging and predictive techniques used by manufacturers to estimate MTTF.

- **Assuming uncorrelated failures.** Evidence from deployed systems suggests that when one fault occurs, other nearby devices are more likely to fail soon. Many factors can cause such correlation. For example, manufacturing irregularities can cause a batch of disks to be substandard, and an organization that purchases disks in bulk may find an entire batch of disks failing at the same time. As another example, disks in the same machine or rack may be of a similar age, may experience similar environmental stress and workloads, and may wear out at a similar time.

- **Confusing a device's MTTF with its useful life.** If a device has an MTTF of one million or more hours, it does not mean that it is expected to last for 100 years or more. Disks are designed to be operated for some finite lifetime, perhaps 5 years. A disk's advertised annual failure rate (i.e., 1/MTTF) applies during the disk's intended service life. As that lifetime is approached, failure rates may rise as the device wears out.

- **Assuming constant failure rates.** A device may have different failure rates over its lifetime. Some devices exhibit *disk infant mortality*, where their failure rate may be higher than normal during their first few weeks of use as latent manufacturing defects are exposed. Others exhibit *disk wear out*, where their failure rate begins to rise after some years.

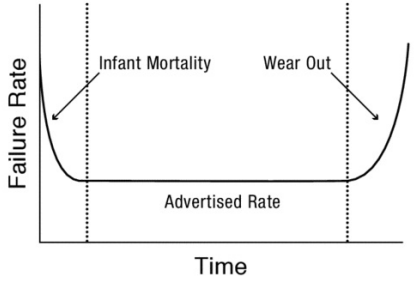


Figure 14.7: Bathtub model of device lifetimes.

A simple model for understanding infant mortality and wear out is the *bathtub model* illustrated in Figure 14.7.

- **Ignoring warning signs.** Some device failures happen without warning, but others are preceded by increasing rates of non-fatal anomalies. Many storage devices implement the *SMART (Self-Monitoring, Analysis, and Reporting Technology)* interface, which provides a way for the operating system to monitor events that may be useful in predicting failures such as read errors, sector remappings, inaccurate seek attempts, or failures to spin up to the target speed.

- **Assuming devices behave identically.** Different device models or even different generations of the same model may have significantly different failure behaviors. One generation might exhibit significantly higher failure rates than expected and the next might exhibit significantly lower rates.

EXAMPLE: Disk failures in large systems. Suppose you have a departmental file server with 100 disks, each with an estimated MTTF of 1.5×10^6 hours. Estimate the expected time until one of those disks fails. For simplicity, assume that each disk has a constant failure rate and that disks fail independently.

ANSWER: If each disk has a MTTF of 1.5×10^6 hours, then 100 disks fail at a 100 times greater rate, giving us a MTTF of 1.5×10^4 hours. So, although the annual failure rate of a single disk is (1 failure / 1.5×10^6 hours) × 24 hours/day × 365 days/year = 0.00585 failures/year, the annual failure rate of the 100 disk system is 0.585 = 58.5%. □

EXAMPLE: Pitfalls. Given the pitfalls discussed above, is this calculation above likely to overestimate or underestimate the failure rate of the system?

ANSWER: Of the factors listed above, the pitfall of *relying on advertised failure rates* seems most significant, and it could lead us to significantly *underestimate* the failure rate of the system.

This solution does *assume constant failure rates*. If the disks are very new or very old, they may suffer higher failure rates than expected, which might cause us to *underestimate* the failure rate of the system.

Because we are only interested in the average rate, the *correlation* pitfall is probably not particularly relevant to our analysis. □

The exponential distribution

When — as in the example — device failures occur at a constant rate, the number of failure events in a fixed time period can be mathematically modeled as a Poisson process, and the interarrival time between failure events follows an exponential distribution.

The exponential distribution is *memoryless* — since the rate of failure events is constant across time, then the expected time to the next failure event is the same — no matter what the current time and no matter how long it has been since the last failure. Thus, if a device has an annual failure rate of 0.5 and thus a mean time to failure of 2 years, and we have been operating the device without a failure for a year, the expected time from the current time to the next failure is still 2 years.

If random variable T represents the time between failures and has an exponential distribution with λ representing the average number of failure events per unit of time, then the probability density function $f_T t$ is:

$$f_T(t) = \lambda e^{-\lambda t} \text{ if } t \geq 0$$

$$= 0 \quad \text{if } t < 0$$

The mean time to failure is MTTF = 1 / λ.

Exponential distributions have a number of convenient mathematical properties. For example, because the failure rate is constant, the mean time to failure is the inverse of the failure rate; this is why it is easy to convert between MTTF and annual failure rates in storage specifications. Also, if the expected number of failures is given for one duration (e.g., 0.1 failures per year), it can easily be converted to the expected number for a different duration (e.g., 0.0003 failures per day). Finally, if we have k independent failure processes with rates of $\lambda_1, \lambda_2, \ldots, \lambda_k$, then the aggregate failure function — the rate at which failures of any of the k kinds occurs — is

$$\lambda_{tot} = \lambda_1 + \lambda_2 + \ldots + \lambda_k$$

and the mean time to the next failure of any kind is $MTTF_{tot} = 1 / \lambda_{tot}$. For example, if we have 100 disks, each with a $MTTF_{disk} = 1.5 \times 10^6$ hours or, equivalently, each failing at a rate of 0.00585 failures per year, then the overall 100-disk system suffers failures at a rate of 100 × 0.00585 = 0.585 failures per year or, equivalently, the 100-disk system has $MTTF_{100disks} = 1.5 \times 10^4$ hours.

Warning. Because the exponential distribution is so mathematically convenient, is tempting to use it even when it is not appropriate. Remember that failures in real systems may be correlated (i.e., they are not independent) and may vary over time (i.e., they are not constant).

14.2.2 RAID: Multi-Disk Redundancy for Error Correction

Given the limits of physical storage devices, storage systems use additional techniques to get acceptable end-to-end reliability. In particular, rather than trying to engineer perfectly reliable (and extremely expensive) storage devices, storage systems use Redundant Arrays of Inexpensive Disks (RAIDs) so that a partial or total failure of one device will not cause data to be lost.

Basic RAIDs

A *Redundant Array of Inexpensive Disks (RAID)* is a system that spreads data redundantly across multiple disks in order to tolerate individual disk failures. Note that the term RAID traditionally refers to redundant *disks*, and for simplicity, we will discuss RAID in the context of disks. The principles, however, apply equally well to other storage devices like flash drives.

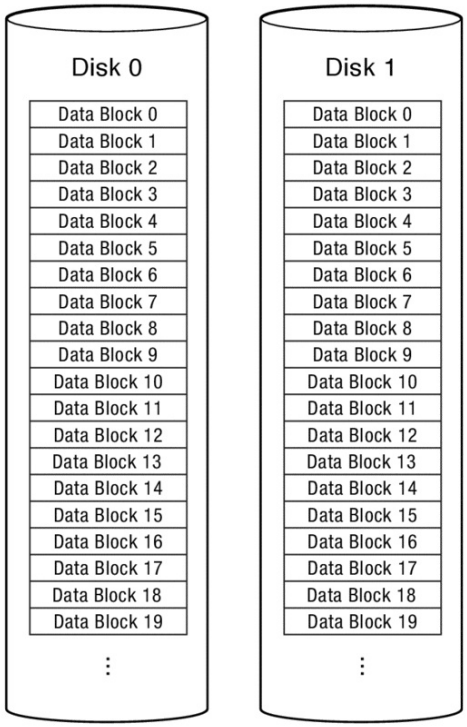


Figure 14.8: RAID 1 with mirroring.

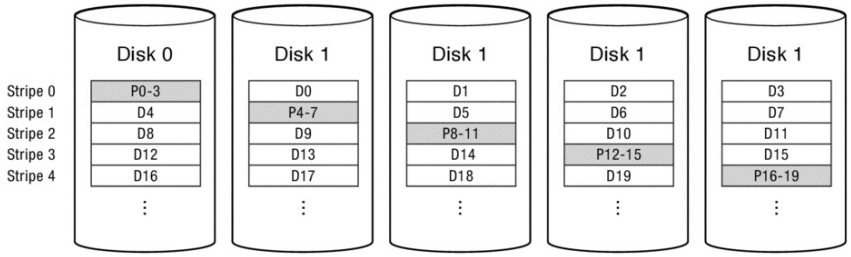


Figure 14.9: RAID 5 with rotating parity.

Figures 14.8 and 14.9 illustrate two common RAID architectures: mirroring and rotating parity.

- **Mirroring.** In RAIDs with *mirroring* (also called *RAID 1*), the system writes each block of data to two disks and can read any block of data from either disk as Figure 14.8 illustrates. If one of the disks suffers a sector or whole-disk failure, no data is lost because the data can still be read from the other disk.
- **Rotating parity.** In RAIDs with *rotating parity* (also called *RAID 5*), the system reduces replication overheads by storing several blocks of data on several disks and protecting those blocks with one redundant block stored on yet another disk as Figure 14.9 illustrates.

  In particular, this approach uses groups of G disks, and writes each of G - 1 blocks of data to a different disk and 1 block of parity to the remaining disk. Each bit of the parity block is produced by computing the exclusive-or of the corresponding bits of the data blocks:

$$\text{parity} = \text{data}_0 \oplus \text{data}_1 \oplus \ldots \oplus \text{data}_{G-1}$$

  If one of the disks suffers a sector or whole-disk failure, lost data blocks can be reconstructed using the corresponding data and parity blocks from the other disks. Note that because the system already knows which disk has failed, parity is sufficient for error correction, not just error detection. For example, if the disk containing block data_0 fails, the block can be reconstructed by computing the exclusive-or of the parity block and the remaining data blocks:

$$\text{data}_0 = \text{parity} \oplus \text{data}_1 \oplus \ldots \oplus \text{data}_{G-1}$$

  To maximize performance, rotating parity RAIDs carefully organize their data layout by rotating parity and striping data to balance parallelism and sequential access:

  - **Rotating parity.** Because the parity for a given set of blocks must be updated each time any of the data blocks are updated, the average parity block tends to be accessed more often than the average data block. To balance load, rather than having G-1 disks store only data blocks and 1 disk store only parity blocks, each disk dedicates 1 / Gth of its space to parity and is responsible for storing 1 / Gth of the parity blocks and (G - 1) / G of the data blocks.
  - **Striping data.** To balance parallelism versus sequential-access efficiency, a *strip* of several sequential blocks is placed on one disk before shifting to another disk for the next strip. A set of G - 1 data strips and their parity strip is called a *stripe*.

    By striping data, requests larger than a block but smaller than a strip require just

one disk to seek and then read or write the full sequential run of data rather than requiring multiple disks to seek and then read smaller sequential runs. Conversely, the RAID can service more widely spaced requests in parallel.

Combining rotating parity and striping, we have the arrangement shown in Figure 14.9.

EXAMPLE: Updating a RAID with rotating parity. For the rotating parity RAID in Figure 14.9, suppose you update data block 21. What disk I/O operations must occur?

ANSWER: The challenge is that we must not only update data block 21, we must also update the corresponding parity block. Since data block 21 is block 1 of its strip and the strip is part of stripe 1, we need to update parity block 1 of the parity strip for stripe 1 (*Parity (1,1,1)* in the figure).

It takes 4 I/O operations to update both the data and parity. First we read the old data D_{21} and parity $P_{1,1,1}$ and "remove" the old data from the parity calculation $P_{tmp} = P_{1,1,1} \oplus D_{21}$. Then we can compute the new parity from the new data $P'_{1,1,1} = P_{tmp} \oplus D'_{21}$. **Finally, we can write the new data D'_{21} and parity $P'_{1,1,1}$** to disks 2 and 1, respectively. □

Atomic update of data and parity

A challenge in implementing RAID is atomically updating both the data and the parity (or both data blocks in a RAID with mirroring).

Consider what would happen if the RAID system in Figure 14.9 crashes in the middle updating block 21, after updating the data block on disk 2 but before updating the parity block on disk 1. Now, if disk 2 fails, the system will reconstruct the wrong (old) data for block 21.

The situation may be even worse if a write to a mirrored RAID is interrupted. Because reads can be serviced by either disk, reads of the inconsistent block may sometimes return the new value and sometimes return the old one.

Solutions. Three solutions and one non-solution are commonly used to solve (or not) the atomic update problem.

- **Non-volatile write buffer.** Hardware RAID systems often include a battery-backed write buffer. An update is removed from the write buffer only once it is safely on disk. The RAID's startup procedures ensure that any data in the write buffer is written to disk after a crash or power outage.
- **Transactional update.** RAID systems can use transactions to atomically update both the data block and the parity block. For example, Oracle's RAID-Z integrates RAID striping with the ZFS file system to avoid overwriting data in place and to atomically update data and parity.
- **Recovery scan.** After a crash, the system can scan all of the blocks in the system and update any inconsistent parity blocks. Note that until that scan is complete, some parity blocks may be inconsistent, and incorrect data may be reconstructed if a disk fails. The Linux md (multiple device) software RAID driver uses this approach.
- **Cross your fingers.** Some software and hardware RAID implementations do not ensure that the data and parity blocks are in sync after a crash. *Caveat emptor.*

RAIDs with rotating parity have high overheads for small writes. Their overheads are far smaller for reads and for full-stripe writes.

RAID levels

An early paper on RAIDs, "A Case for Redundant Arrays of Inexpensive Disks (RAID)" by David Patterson, Garth Gibson, and Randy Katz http://dl.acm.org/citation.cfm?id=50214 described a range of possible RAID organizations and named them RAID 0, RAID 1, RAID 2, RAID 3, RAID 4, and RAID 5. Several of these RAID levels were intended to illustrate key concepts rather than for real-world deployment.

Today, three of these variants are in wide use:

- **RAID 0: JBOD.** RAID level 0 spreads data across multiple disks without redundancy. Any disk failure results in data loss. For this reason, the term RAID is somewhat misleading, and this organization is often referred to as JBOD (Just a Bunch Of Disks).
- **RAID 1: Mirroring.** RAID level 1 mirrors identical data to two disks.
- **RAID 5: Rotating Parity.** RAID level 5 stripes data across G disks. G - 1 of the disks in a stripe store G-1 different blocks of data and the remaining disk stores a parity block. The role of storing the parity block for different data blocks is rotated among the disks to balance load.

Subsequent to the "Case for RAID" paper, new organizations emerged, and many of them were named in the same spirit. Some of these names have become standard.

- **RAID 6: Dual Redundancy.** RAID level 6 is similar to RAID level 5, but instead of one parity block per group, two redundant blocks are stored. These blocks are generated using erasure codes such as Reed-Solomon codes that allow reconstruction of all of the original data as long as at most two disks fail.
- **RAID 10 and RAID 50: Nested RAID.** RAID 10 and RAID 50 were originally called RAID 1+0 and RAID 5+0. They simply combine RAID 0 with RAID 1 or RAID 5. For example, a RAID 10 system mirrors pairs of disks for redundancy (RAID 1), treats each pair of mirrored disks as a single reliable logical disk, and then stripes data non-redundantly across these logical disks (RAID 0).

Many other RAID levels have been proposed. In some cases, these new "levels" have more to do with marketing than technology. ("Our company's RAID 99+ is much better than your company's puny RAID 14.") In any event, we regard the particular nomenclature used to describe exotic RAID organizations as relatively unimportant; our

discussion focuses on mirroring (RAID 1), rotating parity (RAID 5), and dual redundancy (RAID 6). Other organizations can be analyzed using principles from these approaches.

Recovery. In either RAID arrangement, if a disk suffers a sector failure, the disk reports an error when there is an attempt to read the sector and, if necessary, remaps the damaged sector to a spare one. Then, the RAID system reconstructs the lost sector from the other disk(s) and rewrites it to the original disk.

If a disk suffers a whole-disk failure, an operator replaces the failed disk, and the RAID system reconstructs all of the disk's data from the other disk(s) and rewrites the data to the replacement disk. The average time from when a disk fails until it has been replaced and rewritten is called the *mean time to repair (MTTR)*.

RAID Reliability

A RAID with one redundant disk per group (e.g., mirroring or rotating parity RAIDs) can lose data in three ways: two full disk failures, a full disk failure and one or more sector failures on other disks, and overlapping sector failures on multiple disks. The expected time until one of these events occurs is called the *mean time to data loss (MTTDL)*.

Two full-disk failures. If two disks fail, the system will be unable to reconstruct the missing data.

To get a sense of how serious a problem this might be, suppose that a system has N disks with one parity block per G blocks, and suppose that disks fail independently with a mean time to failure of MTTF and a mean time to replace a failed disk and recover its data of MTTR.

Then, when the system is operating properly, the expected time until the first failure is MTTF / N. Assuming MTTR << MTTF , there is essentially a race to replace the disk and reconstruct its data before a second disk fails. We lose this race and hit the second failure before the repair is done with probability MTTF / ((G - 1) × MTTR), giving us a mean time to data loss from multiple full-disk failures of

$$\text{MTTDL}_{\text{two-full-disk}} = \text{MTTF}^2 / (\text{N} \times (\text{G} - 1) \times \text{MTTR})$$

EXAMPLE: Mean time to double-disk failure. Suppose you have 100 disks organized into groups of 10, with one disk storing a parity block per nine disks storing data blocks. Assuming that disk failures are independent and the per-disk mean time to failure is 10^6 hours and assuming that the mean time to repair a failed disk is 10 hours, estimate the expected mean time to data loss due to a double-disk failure.

ANSWER: Because failures are assumed to occur independently and at a constant rate, we can use the equation above:

$$\begin{aligned}\text{MTTDL}_{\text{two-full-disk}} &= \text{MTTF}^2 / (\text{N} \times (\text{G} - 1) \times \text{MTTR}) \\ &= (10^6 \text{ hours})^2 / (10^2 \times 9 \times 10 \text{ hours}) \\ &\approx \mathbf{10^8 \text{ hours}}\end{aligned}$$

Thus, assuming independent failures at the expected rate and assuming no other sources of data loss, this organization appears to have raised the mean time to data loss from about 100 years (for a single disk) to about 10,000 years (for 90 disks worth of data and 10 disks worth of parity). □

One full-disk failure and a sector failure. If one or more disks suffer sector failures and another disk suffers a full-disk failure, the RAID system cannot recover all of its data. Assuming independent failures that arrive at a constant rate, we can estimate probability of data loss over some interval as the probability of suffering a disk failure times the probability that we will fail to read all data needed to reconstruct the lost disk's data:

$$\begin{aligned}\text{P}_{\text{lostDataFromDiskAndSector}} &= \text{P}_{\text{DiskFailure}} \times \text{P}_{\text{RecoveryError}} \\ &= (\text{N} / \text{MTTF}) \times \text{P}_{\text{fail\_recovery\_read}}\end{aligned}$$

If this gives us the probability of losing data over some period of time or equivalently the rate of data-loss failures, then inverting this equation gives us the mean time to data loss (MTTDL). Thus, we can estimate the mean time to data loss from this failure mode based on the expected time between full disk failures divided by the odds of failing to read all data needed to reconstruct the lost disk's data.

$$\text{MTTDL}_{\text{disk+sector}} = (\text{MTTF} / \text{N}) \times (1 / \text{P}_{\text{fail\_recovery\_read}})$$

EXAMPLE: Mean time to failed disk and failed sector. Assuming that during recovery, latent sector errors are discovered at a rate of 1 per 10^{15} bits read and assuming that the mean time to failure for each of 100 1 TB disks organized into groups of 10 is 10^6 hours,

what is the expected mean time to data loss due to full-disk failure combined with a sector failure?

ANSWER:

$$\text{MTTDL}_{\text{disk+sector}} = (\text{MTTF} / N) \times (1 / P_{\text{fail\_recovery\_read}})$$

$$= (10^6 / 100) \times (1 / P_{\text{fail\_recovery\_read}})$$

To estimate $P_{\text{fail\_recovery\_read}}$ we will assume that each bit fails independently and is successfully read with probability $(1/(1 - 10^{-15}))$. Then the probability of reading 1 TB from each of 9 disks is:

$$\begin{aligned} P_{succeed\_recovery\_read} &= P_{succeed\_bit\_read}^{number of bits} \\ &= (1-10^{-15})^{9\ \text{disks} \times 10^{12} \frac{\text{bytes}}{\text{disk}} \times 8 \frac{\text{bits}}{\text{byte}}} \\ &\approx 0.93 \end{aligned}$$

So, there is roughly a 93% chance that recovery will succeed and a 7% chance that recovery will fail. We then have

$$\text{MTTDL}_{\text{disk+sector}} = (10^6 / 100) \times (1 / .07)$$

$$= \mathbf{1.4 \times 10^5\ hours}$$

Notice that this rate of data loss is much higher than the rate from double disk failures calculated above. Of course, the relative contributions of each failure mode will depend on disks' MTTF, size, and bit error rates as well as the system's MTTR. □

Failure of two sectors sharing a redundant sector. In principle, it is also possible to lose data because the corresponding sectors fail on different disks. However, with billions of distinct sectors on each disk and small numbers of latent failures per disk, this failure mode is likely to be a negligible risk for most systems.

Overall data loss rate. If we assume independent failures and constant failure rates, then we can add the failure rates from the two significant failure modes to estimate the combined failure rate:

$$\begin{aligned} FailureRate_{indep+const} &= FailureRate_{two-full-disk} + FailureRate_{disk+sector} \\ &= \frac{1}{MTTDL_{two-full-disk}} + \frac{1}{MTTDL_{disk+sector}} \\ &= \frac{N(G-1)MTTR}{MTTF^2} + \frac{N \times P_{fail\_recovery\_read}}{MTTF} \\ &= \frac{N}{MTTF}\left(\frac{MTTR(G-1)}{MTTF} + P_{fail\_recovery\_read}\right) \end{aligned}$$

The total failure rate is thus the rate that the first disk fails times the rate that either a second disk in the group fails before the repair is completed or a sector error is encountered when the disks are being read to rebuild the lost disk.

We label the above $\text{FailureRate}_{\text{indep+const}}$ to emphasize the strong assumptions of independent failures and constant failure rates. As noted above, failures are likely to be correlated in many environments and failure rates of some devices may increase over time. Both of these factors may result significantly higher failure rates than expected.

EXAMPLE: Combined failure rate. For the system described in the previous examples (100 disks, rotating parity with a group size of 10, mean time to failure of 10^6 hours, mean time to repair of 10 hours, and non-recoverable read error rate of one sector per 10^{15} bits) assuming that all failures are independent, estimate the MTTDL when both double-disk and single-disk-and-sector failures are considered.

ANSWER:

$$\begin{aligned} FailureRate_{indep+const} &= \frac{N}{MTTF}\left(\frac{MTTR(G-1)}{MTTF} + P_{fail\_recovery\_read}\right) \\ &= \frac{100\ \text{disks}}{10^6}\ \text{hours}\left(\frac{10\ \text{hours}}{10^6\ \text{hours}} + 0.0694\right) \\ &= \frac{1}{10^4\ \text{hours}}\left(\frac{1}{10^4} + 0.0694\right) \\ &= \frac{1}{10^4\ \text{hours}}(0.0695) \\ &= 6.95 \times 10^{-6} \frac{\text{failures}}{\text{hour}} \end{aligned}$$

Inverting the failure rate gives the mean time to data loss:

$$\text{MTTDL}_{\text{const+indep}} = \frac{1}{\text{FailureRate}_{\text{indep+const}}}$$

$$= 1.44 \times 10^5\ \text{hours/failure}$$

$$= \mathbf{16.4}\ \text{years/failure}$$

□

Two things in the example above are worth special note. First, for these parameters, the

dominant cause of data loss is likely to be a single disk failure combined with a non-recoverable read error during recovery. Second, for these parameters and this configuration, the resulting 6% chance of losing data per year may be unacceptable for many environments. As a result, systems use various techniques to improve the MTTDL in RAID systems.

Improving RAID Reliability

What can be done to further improve reliability? Broadly speaking, we can do three things: (1) increase redundancy, (2) reduce non-recoverable read error rates, and (3) reduce mean time to repair. All of these approaches, in various combinations, are used in practice.

Here are some common approaches:

Increasing redundancy with more redundant disks. Rather than having a single redundant block per group (e.g., using two mirrored disks or using one parity disk for each stripe) systems can use double redundancy (e.g., three disk replicas or two error correction disks for each stripe). In some cases, systems may use even more redundancy. For example, the Google File System (GFS) is designed to provide highly reliable and available storage across thousands of disks; by default, GFS stores each data block on three different disks.

A *dual redundancy array* ensures that data can be reconstructed despite any two failures in a stripe by generating two redundant blocks using erasure codes such as Reed-Solomon codes. This approach is sometimes called *RAID 6*.

A system with dual redundancy can be much more reliable than a simple single redundancy RAID. With dual redundancy, the most likely data loss scenarios are (a) three full-disk failures or (b) a double-disk failure combined with one or more non-recoverable read errors.

If we optimistically assume that failures are independent and occur at a constant rate, a system with two redundant disks per stripe has a potentially low combined data loss rate:

$$\begin{aligned} FailureRate_{dual+indep+const} &= \frac{N}{MTTF} \\ &\times \frac{MTTR(G-1)}{MTTF} \\ &\times (\frac{MTTR(G-2)}{MTTF} + P_{fail\_recovery\_read}) \end{aligned}$$

This data loss rate is nearly MTTF / (MTTR × (G - 1)) times better than the single-parity data loss rate; for disks with MTTFs of over one million hours, MTTRs of less than 10 hours, and group sizes of ten or fewer disks, double parity improves the estimated rate by about a factor of 10,000.

We emphasize, however, that the above equation almost certainly underestimates the likely data loss rate for real systems, which may suffer correlated failures, varying failure rates, higher failure rates than advertised, and so on.

Reducing non-recoverable read errors with scrubbing. A storage device's sector-level error rate is typically expressed as a single *non-recoverable read rate*, suggesting that the rate is constant. The reality is more complex. Depending on the device, errors may accumulate over time and heavier workloads may increase the rate at which errors accumulate.

An important technique for reducing a disk's non-recoverable read rate is *scrubbing*: periodically reading the entire contents of a disk, detecting sectors with unrecoverable read errors, reconstructing the lost data from the remaining disks in the RAID array, and attempting to write and read the reconstructed data to and from the suspect sector. If writes and reads succeed, then the error was caused by a transient fault, and the disk continues to use the sector, but if the sector cannot be successfully accessed, the error is permanent, and the system remaps that sector to a spare and writes the reconstructed data there.

Reducing non-recoverable read error rates with more reliable disks. Different disk models promise significantly different non-recoverable read error rates. In particular, in 2011, many disks aimed at laptops and personal computers claim unrecoverable read error rates of one per 10^{14} bits read, while disks aimed at enterprise servers often have lower storage densities but can promise unrecoverable read error rates of one per 10^{16} bits read. This two order of magnitude improvement greatly reduces the probability that a RAID system loses data from a combination of a full disk failure and a non-recoverable read error during recovery.

Reducing mean time to repair with hot spares. Some systems include "hot spare" disk drives that are idle, but plugged into a server so that if one of the server's disks fails, the hot spare can be automatically activated to replace the lost disk.

Note that even with hot spares, the mean time to repair a disk is limited by the time it takes to write the reconstructed data to it, and this time is often measured in hours. For example, if we have a 1 TB disk and can write at 100 MB/s, the mean time to repair for the disk will be at least 10^4 seconds — about 3 hours. In practice, repair time may be even larger if the bandwidth achieved is less than assumed here.

Reducing mean time to repair with declustering. Disks with hundreds of gigabytes to a few terabytes can take hours to fully write with reconstructed data. *Declustering* splits reconstruction of a failed disk across multiple disks. Declustering thus allows parallel reconstruction, thus speeding up reconstruction and reducing MTTR.

For example, the Hadoop File System (HDFS) is a cluster file system that writes each data block to three out of potentially hundreds or thousands of disks. It chooses the three disks for each block more or less randomly. If one disk fails, it re-replicates the lost blocks approximately randomly across the remaining disks. If we have N disks each with a bandwidth of B, total reconstruction bandwidth can approach (N / 2) × B; for example, if there are 1000 disks with 100 MB/s bandwidths, reconstruction bandwidth can theoretically approach 500 GB/s, allowing re-replication of a 1 TB disk's data in a few seconds.

In practice, re-replication will be slower than this for at least three reasons. First, resources other than the disk (e.g., the network) may bottleneck recovery. Second, the system may throttle recovery speed to avoid starving user requests. Third, if a server crashes and its disks become inaccessible, the system may delay starting recovery — hoping that the

server will soon recover — to avoid imposing extra load on the system.

Pitfalls

When constructing a reliable storage system, it is not enough to plug provide enough redundancy to tolerate a target number of failures. We also need to consider how failures are likely to occur (e.g., they may be correlated) and what it takes to correct them (e.g., successfully reading a lot of other data). More specifically, be aware of the following pitfalls:

- **Assuming uncorrelated failures..** It is easy to get gaudy MTTDL numbers by adding a redundant device or two and multiplying the devices' MTTFs. But the simple equation for MTTDL we derived above only applies when failures are uncorrelated. Even a 1% chance of correlated failures dramatically changes the estimate. Unfortunately, it is often difficult to estimate correlation rates *a priori*, so designers must sometimes just add a significant safety margin and hope that it is enough.
- **Ignoring the risk from latent errors..** It is not uncommon to see analyses of RAID reliability that considers full device failures but not non-recoverable read failures. As we have seen above, non-recoverable read errors can dramatically reduce the probability of successfully recovering data after a disk failure.
- **Not implementing scrubbing..** Periodically scrubbing disks to detect and correct latent errors can significantly reduce the risk of data loss. Although it can be difficult to predict the appropriate scrubbing frequency *a priori*, a system that uses scrubbing can monitor the rate at which non-correctable read errors are found and corrected and use the measured rate to adjust the scrubbing frequency.
- **Not having a backup..** The techniques discussed in this section can protect a system against many, but not all, faults. For example, a widespread correlated failure (e.g., a building burning down), an operator error (e.g., "rm -r *"), or a software bug could corrupt or delete data stored across any number of redundant devices.

  A *backup* system provides storage that is separate from a system's main storage. Ideally, the separation is both physical and logical.

  Physical separation means that backup storage devices are in different locations than the primary storage devices. For example, some systems achieve physical separation by copying data to tape and storing the tapes in a different building than the main storage servers. Other systems achieve physical separation by storing data to remote disk arrays such as those provided by cloud backup and disaster recovery services.

  Logical separation means that the interface to the backup system is restricted to prevent premature deletion of data. For example, some backup systems provide an interface that allows a user to read *but not write* old versions of a file (e.g., the file as it existed one hour, two hours, four hours, one day, one week, one month, and one year ago).

Modeling real systems

The equations in the main text for estimating a system's mean time to data loss are only applicable if failure rates are constant and if failures are uncorrelated. Unfortunately, empirical studies often observe correlation among full-disk failures, among sector-level failures, and between sector-level and full-disk failures, and they frequently find failure rates that vary significantly with disks' ages. Unfortunately, if failure rates vary over time or failures are correlated, the failure arrival distribution is no longer described by an exponential distribution, and the math quickly gets difficult.

One solution is to use randomized simulation to estimate the probability of data loss over some duration of interest. For example, we might want to estimate the probability of losing data over 10 years for a 1000-disk system organized in groups of 10 disks with rotating parity.

To do this, our simulation would track which disks are functioning normally, which have latent sector errors, and which have suffered full disk failures. The transitions between states could be based on measurement studies or field data on key factors: (a) the rate that disks suffer full disk failures (possibly dependent on the disks' ages, the number of recent full disk failures, or the number of individual sector failures a disk has had); (b) the rate at which sector failures arise (possibly dependent on the age of the disk, workload of the disk, and recent frequency of sector failures); (c) the repair time when a disk fails; and (d) the expected time for scrubbing to detect and repair a sector error.

To estimate the probability of data loss, we would repeatedly simulate the system for a decade and count the number of times the system enters a state in which data is lost (i.e., a group has two full disk failures or has both a full disk failure and a sector failure on another disk).

14.2.3 Software Integrity Checks

Although storage devices include sector- or page-level checksums to detect data corruption, many recent file systems have included additional, higher-level, checksums and other integrity checks on their data.

These checks can catch a range of errors that hardware-level checksums can miss. For example, they can detect *wild writes* or *lost writes* where a bug in the operating system software, device driver software, or device firmware misdirects a write to the wrong block or page or fails to complete an intended write. They can also detect rare *ECC false negatives* when the hardware-level error correcting codes fail to detect a multi-bit corruption.

When a software integrity check fails on a block read or during latent-error scrubbing, the system reconstructs the lost or corrupted block using the redundant storage in the RAID.

Two examples of software integrity checks used today are *block integrity metadata* and *file system fingerprints*.

Block integrity metadata. Some file systems, like Network Appliance's WAFL file

system, include *block integrity metadata* that allows the software to validate the results of each block it reads.

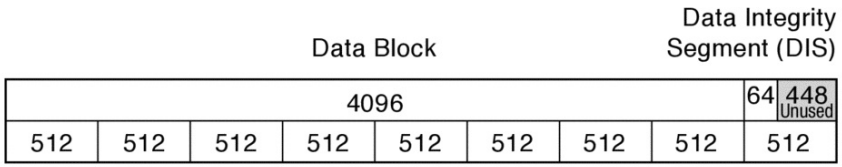
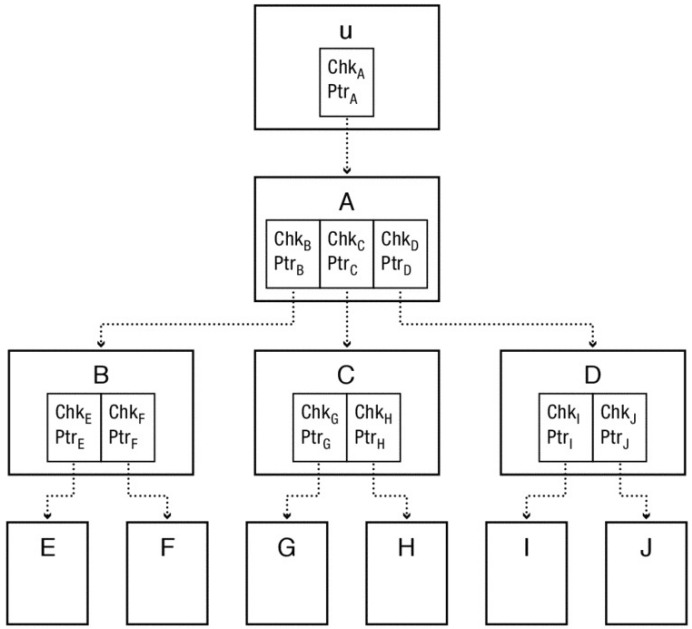


Figure 14.10: To improve reliability Network Appliance's WAFL file system stores a 64 byte data integrity segment (DIS) with each 4 KB data block.

As Figure 14.10 illustrates, WAFL stores a 64 byte *data integrity segment* (DIS) with each 4 KB data block. The DIS contains a checksum of the data block, the identity of the data block (e.g., the ID of the file to which it belongs and the block's offset in that file), and a checksum of the DIS, itself.

Then, when a block is read, the system performs three checks. First, it checks the DIS's checksum. Second, it verifies that the data in the block corresponds to the checksum in the block's data integrity segment. Third, it verifies that the identity in the block's DIS corresponds to the file block it was intending to read. If all of these checks pass, the file system can be confident it is returning the correct data; if not, the file system can reconstruct the necessary data from redundant disks in the RAID.

File system fingerprints. Some file systems, like Oracle's ZFS, include *file system fingerprints* that provide a checksum across the entire file system in a way that allows efficient checks and updates when individual blocks are read and written.



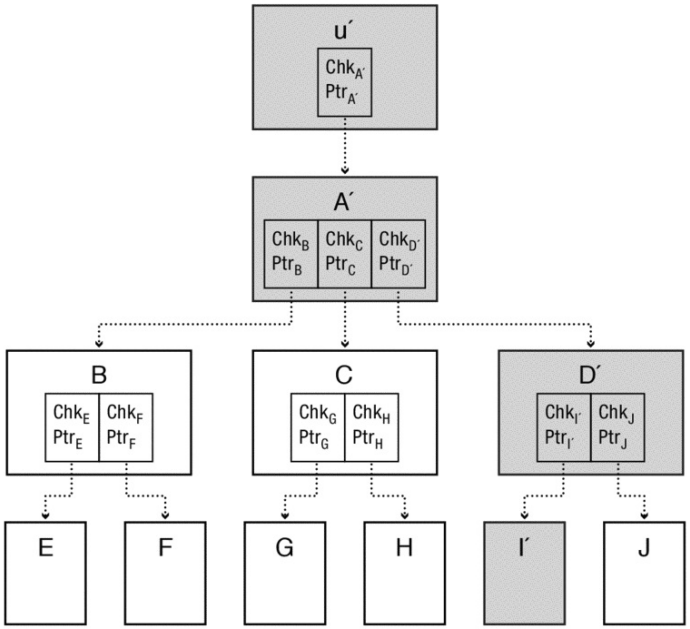


Figure 14.11: ZFS stores all data in a Merkle tree so that each node of the tree includes both a pointer to and a checksum of each of its children (Chk and Ptr in the figure). On an update, all nodes from the updated block (I') to the root (u') are changed to reflect the new pointer and checksum values.

As illustrated in Figure 14.11, all of ZFS's data structures are arranged in a tree of blocks with a root node called the uberblock. At each internal node of the tree, each reference to a child node includes both a pointer to and a checksum of the child. Thus, the reference to any subtree includes a checksum that covers all of that subtree's contents, and the uberblock holds a checksum that covers the entire file system.

When ZFS reads data (i.e., leaves of the tree) or metadata (i.e., internal nodes of the tree), it follows the pointers down the tree to find the right block to read, computing a checksum of each internal or leaf block and comparing it to the checksum stored with the block reference. Similarly, as Figure 14.11 illustrates, when ZFS writes a block, it updates the references from the updated block to the uberblock so that each includes both the new checksum and (since ZFS never updates data structures in place) new block pointer.

Layers upon layers upon layers In this chapter we focus on error detection and correction at three levels: the individual storage devices (e.g., disks and flash), storage architectures (e.g., RAID), and file systems.

Today, storage systems with important data often include not just these layers, but additional ones. Enterprise and cloud storage systems distribute data across several geographically distributed sites and may include high-level checksums on that geographically replicated data. Within a site, they may replicate data across multiple servers using what is effectively a distributed file system. At each server, the distributed file system may store data using a local file system that includes file-system-level checksums on the locally stored data. And, invariably, the local server will use storage devices that detect and sometimes correct low-level errors.

Although we do not discuss cross-machine and geographic replication in any detail, the principles described in this chapter also apply to these systems.

14.3 Summary and Future Directions

Although individual storage devices include internal error correcting codes, additional redundancy for error detection and correction is often needed to provide acceptably reliable storage. In fact, today, it is seldom acceptable to store valuable data on a single device without some form of RAID-style redundancy. By the same token, many if not most file systems designed over the past decade have included software error checking to catch data corruption and loss occurrences that are not detectable by device-level hardware checks.

Increasingly now and in the future, systems go beyond just replicating data across multiple disks on a single server to distributed replication across multiple servers. Sometimes these replicas are configured to protect data even if significant physical disasters occur.

For example, Amazon's Simple Storage Service (S3) allows customers to pay a monthly fee to store data on servers run by Amazon. Amazon states that the system is "designed to provide 99.999999999% durability of objects over a given year." To provide such high reliability, S3 stores data at multiple data centers, quickly detects and repairs lost redundancy, and validates checksums of stored data.

Exercises

1. Suppose that a text editor application uses the rename technique for safely saving updates by saving the updated file to a new filed (e.g., #doc.txt# and then calling rename("#doc.txt#", "doc.txt") to change the name of the updated file from #doc.txt# to doc.txt. POSIX rename promises that the update to doc.txt will be atomic — even if a crash occurs, doc.txt will refer to either the old file or the new one. However, POSIX does not guarantee that the entire rename operation will be atomic. In particular, POSIX allows implementations in which there is a window in which a crash could result in a state where both doc.txt and #doc.txt# refer to the same, new document.
    a. How should a text-editing application react if, on startup, it sees both doc.txt and doc.txt and (i) both refer the same file or (ii) each refers to a file with different contents?
    b. Why might POSIX permit this corner case (where we may end up with two names that refer to the same file) to exist?

c. Explain how an FFS-based file system without transactions could use the "ad hoc" approach discussed in Section 14.1.1 to ensure that (i) doc.txt always refers to either the old or new file, (ii) the new file is never lost – it is always available as at least one of doc.txt or #doc.txt#, and (iii) there is some window where the new file may be accessed as both doc.txt and #doc.txt#.

d. Section 14.1.1 discusses three reasons that few modern file systems use the "ad-hoc" approach. However, many text editors still do something like this. Why have the three issues had less effect on applications like text editors than on file systems?

2. Above, we defined two-phase locking for basic mutual exclusion locks. Extend the definition of two-phase locking for systems that use readers-writers locks.

3. Suppose that x and y represent the number of hours two managers have assigned you to work on each of two tasks with a constraint that $x + y \leq 40$. Earlier, we showed that snapshot isolation could allow one transaction to update x and another concurrent transaction to update y in a way that would violate the constraint $x + y \leq 40$. Is such an anomaly possible under serializability? Why or why not?

4. Suppose you have transactional storage system tStore that allows you to read and write fixed-sized 2048-byte blocks of data within transactions, and you run the code in Figure 14.12.

```
...
byte b1[2048];  byte b2[2048];
byte b3[2048];  byte b4[2048];

TransID t1 = tStore.beginTransaction();
TransID t2 = tStore.beginTransaction();
TransID t3 = tStore.beginTransaction();
TransID t4 = tStore.beginTransaction();

// Interface is
//     writeBlock(TransID tid, int blockNum, byte buffer[]);
tStore.writeBlock(t1, 1, ALL_ONES);
tStore.writeBlock(t1, 2, ALL_TWOS);
tStore.writeBlock(t2, 3, ALL_THREES);
tStore.writeBlock(t1, 3, ALL_FOURS);
tStore.writeBlock(t1, 2, ALL_FIVES);
tStore.writeBlock(t3, 2, ALL_SIXES);
tStore.writeBlock(t4, 4, ALL_SEVENS);
tStore.readBlock(t2, 1,  b1);
tStore.commit(t3);
tStore.readBlock(t2, 3,  b2);
tStore.commit(t2);
tStore.readBlock(t1, 3,  b3);
tStore.readBlock(t4, 3,  b4);
tStore.commit(t1);

// At this point, the system crashes
```

Figure 14.12: Sample code for a transactional storage system.

The system crashes at the point indicated above.

a. Assume that ALL_ONES, ALL_TWOS, etc. are each arrays of 2048 bytes with the indicated value. Assume that when the program is started, all blocks in the tStore have the value ALL_ZEROS.

   Just before the system crashes, what is the value of b1 and what is the value of b2?

b. In the program above, just before the system crashes, what is the value of b3 and what is the value of b4?

c. Suppose that after the program above runs and crashes at the indicated point. After the system restarts and completes recovery and all write-backs, what are the values stored in each of blocks 1, 2, 3, 4, and 5 of the tStore?

5. Go to an on-line site that sells hard disk drives, and find the largest capacity disk you can buy for less than $200. Now, track down the spec sheet for the disk and, given the disk's specified bit error rate (or unrecoverable read rate), estimate the probability of encountering an error if you read every sector on the disk once.

6. Suppose we define a RAID's *access cost* as the number disk accesses divided by the number of data blocks read or written. For each of following configurations and workloads, what is the access cost?

a. Workload: a series of random 1-block writes
   Configuration: mirroring

b. Workload: a series of random 1-block writes
   Configuration: distributed parity

c. Workload: a series of random 1-block reads
   Configuration: mirroring

d. Workload: a series of random 1-block reads
   Configuration: distributed parity

e. Workload: a series of random 1-block reads
   Configuration: distributed parity with group size G and one failed disk

f. Workload: a long sequential write
   Configuration: mirroring

g. Workload: a long sequential write
   Configuration: distributed parity with a group size of G

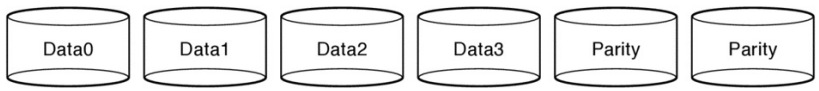


Figure 14.13: Example of a poor design choice for the content of redundant blocks.

7. Suppose that an engineer who has not taken this class tries to create a disk array with dual-redundancy but instead of using an appropriate error correcting code such as Reed-Solomon, the engineer simply stores a copy of each parity block on two disks, as in Figure 14.13.

   Give an example of how a two-disk failure can cause a stripe to lose data in such a system. Explain why data cannot be reconstructed in that case.

8. Some RAID systems improve reliability with intra-disk redundancy to protect against non-recoverable read failures. For example, each individual disk on such a system might reserve one 4KB parity block in every 32 KB extent and then store 28KB (7 4KB blocks) of data and 4 KB (1 4KB block) of parity in each extent.

   In this arrangement, each data block is protected by two parity blocks: one interdisk parity block on a different disk and one intradisk parity block on the same disk.

   This approach may reduce a disk's effective non-recoverable read error rate because if one block in an extent is lost, it can be recovered from the remaining sectors and parity on the disk. Of course, if multiple blocks in the same extent are lost, the system must rely on redundancy from other disks.

   a. Assuming that a disk's non-recoverable read errors are independent and occur at a rate of one lost 512 byte sector per 10^{15} bits read, what is the effective non-recoverable read error rate if the operating system stores one parity block per seven data blocks on the disk?

      Hint: You may find the bc or dc arbitrary-precision calculators useful. These programs are standard in many Unix, Linux, and OSX distributions. See the man pages for instructions.

   b. Why is the above likely to significantly overstate the impact of intra-disk redundancy?

9. Many RAID implementations allow on-line repair in which the system continues to operate after a disk failure, while a new empty disk is inserted to replace the failed disk, and while regenerating and copying data to the new disk.

   Sketch a design for a 2-disk, mirrored RAID that allows the system to remain on-line during reconstruction, while still ensuring that when the data copying is done, the new disk is properly reconstructed (i.e., it is an exact copy of other disk.)

   In particular, specify (1) what is done by a recovery thread, (2) what is done on a read during recovery, and (3) what is done on a write during recovery. Also explain why your system will operate correctly even if a crash occurs in the middle of reconstruction.

10. Suppose you are willing to sacrifice no more than 1% of a disk's bandwidth to scrubbing. What is maximum frequency at which you could scrub a 1 TB disk with 100 MB/s bandwidth?

11. Suppose a 3 TB disk in a mirrored RAID system crashes. Assuming the disks used in the system can sustain 100MB/s sequential bandwidth, what is the minimum mean time to repair that can be achieved? Why might a system be configured to perform recovery slower than this?

Size	
Platters/Heads	2/4
Capacity	320 GB
Performance	
Spindle speed	7200 RPM
Average seek time read/write	10.5 ms/ 12.0 ms
Maximum seek time	19 ms
Track-to-track seek time	1 ms
Transfer rate (surface to buffer)	54–128 MB/s
Transfer rate (buffer to host)	375 MB/s
Buffer memory	16 MB
Reliability	
Nonrecoverable read errors per sectors read	1 sector per 10^{14}
MTBF	600,000 hours
Product life	5 years or 20,000 power-on hours
Power	
Typical	16.35 W
Idle	11.68 W

Figure 14.14: Disk specification

12. Suppose I have a disk such as the one described in Figure 14.14 and a workload consisting of a continuous stream of updates to random blocks of the disk.

Assume that the disk scheduler uses the SCAN/Elevator algorithm.

a. What is the throughput in number of requests per second if the application issues one request at a time and waits until the block is safely stored on disk before issuing the next request?

b. What is the throughput in number of requests per second if the application buffers 100 MB of writes, issues those 100 MB worth of writes to disk as a batch, and waits until those writes are safely on disk before issuing the next 100 MB batch of requests?

   Suppose that we must ensure that – even in the event of a crash – the ith update can be observed by a read after crash recovery only if all updates that preceded the ith update can be read after the crash. That is, we have a FIFO property for updates – the i+1'st update cannot "finish" until the ith update finishes. (1) Design an approach to get good performance for this workload. (2) Explain why your design ensures FIFO even if crashes occur. (3) Estimate your approach's throughput in number of requests per second. (For comparison with the previous part of the problem, your solution should not require significantly more than 100MB of main-memory buffer space.)

c. Design an approach to get good performance for this workload. (Be sure to explain how writes, reads, and crash recovery work.)

d. Explain why your design ensures FIFO even if crashes occur.

e. Estimate your approach's throughput in number of requests per second.