CHAPTER 19

Distributed Databases

Unlike parallel systems, in which the processors are tightly coupled and constitute a single database system, a distributed database system consists of loosely coupled sites that share no physical components. Furthermore, the database systems that run on each site may have a substantial degree of mutual independence. We discussed the basic structure of distributed systems in Chapter 18.

Each site may participate in the execution of transactions that access data at one site, or several sites. The main difference between centralized and distributed database systems is that, in the former, the data reside in one single location, whereas in the latter, the data reside in several locations. This distribution of data is the cause of many difficulties in transaction processing and query processing. In this chapter, we address these difficulties.

We start by classifying distributed databases as homogeneous or heterogeneous, in Section 19.1. We then address the question of how to store data in a distributed database in Section 19.2. In Section 19.3, we outline a model for transaction processing in a distributed database. In Section 19.4, we describe how to implement atomic transactions in a distributed database by using special commit protocols. In Section 19.5, we describe concurrency control in distributed databases. In Section 19.6, we outline how to provide high availability in a distributed database by exploiting replication, so the system can continue processing transactions even when there is a failure. We address query processing in distributed databases in Section 19.7. In Section 19.8, we outline issues in handling heterogeneous databases. In Section 19.9, we describe directory systems, which can be viewed as a specialized form of distributed databases.

19.1 Homogeneous and Heterogeneous Databases

In a **homogeneous distributed database**, all sites have identical database management system software, are aware of one another, and agree to cooperate in processing users' requests. In such a system, local sites surrender a portion of their autonomy

in terms of their right to change schemas or database management system software. That software must also cooperate with other sites in exchanging information about transactions, to make transaction processing possible across multiple sites.

In contrast, in a **heterogeneous distributed database**, different sites may use different schemas, and different database management system software. The sites may not be aware of one another, and they may provide only limited facilities for cooperation in transaction processing. The differences in schemas are often a major problem for query processing, while the divergence in software becomes a hindrance for processing transactions that access multiple sites.

In this chapter, we concentrate on homogeneous distributed databases. However, in Section 19.8 we briefly discuss query processing issues in heterogeneous distributed database systems. Transaction processing issues in such systems are covered later, in Section 24.6.

19.2 Distributed Data Storage

Consider a relation r that is to be stored in the database. There are two approaches to storing this relation in the distributed database:

- **Replication**. The system maintains several identical replicas (copies) of the relation, and stores each replica at a different site. The alternative to replication is to store only one copy of relation r.
- **Fragmentation**. The system partitions the relation into several fragments, and stores each fragment at a different site.

Fragmentation and replication can be combined: A relation can be partitioned into several fragments and there may be several replicas of each fragment. In the following subsections, we elaborate on each of these techniques.

19.2.1 Data Replication

If relation r is replicated, a copy of relation r is stored in two or more sites. In the most extreme case, we have **full replication**, in which a copy is stored in every site in the system.

There are a number of advantages and disadvantages to replication.

- **Availability**. If one of the sites containing relation r fails, then the relation r can be found in another site. Thus, the system can continue to process queries involving r, despite the failure of one site.
- **Increased parallelism**. In the case where the majority of accesses to the relation r result in only the reading of the relation, then several sites can process queries involving r in parallel. The more replicas of r there are, the greater the chance that the needed data will be found in the site where the transaction is executing. Hence, data replication minimizes movement of data between sites.

- **Increased overhead on update**. The system must ensure that all replicas of a relation *r* are consistent; otherwise, erroneous computations may result. Thus, whenever *r* is updated, the update must be propagated to all sites containing replicas. The result is increased overhead. For example, in a banking system, where account information is replicated in various sites, it is necessary to ensure that the balance in a particular account agrees in all sites.

In general, replication enhances the performance of read operations and increases the availability of data to read-only transactions. However, update transactions incur greater overhead. Controlling concurrent updates by several transactions to replicated data is more complex than in centralized systems, which we saw in Chapter 16. We can simplify the management of replicas of relation *r* by choosing one of them as the **primary copy** of *r*. For example, in a banking system, an account can be associated with the site in which the account has been opened. Similarly, in an airline-reservation system, a flight can be associated with the site at which the flight originates. We shall examine the primary copy scheme and other options for distributed concurrency control in Section 19.5.

19.2.2 Data Fragmentation

If relation *r* is fragmented, *r* is divided into a number of *fragments* $r_1, r_2, \ldots, r_n$. These fragments contain sufficient information to allow reconstruction of the original relation *r*. There are two different schemes for fragmenting a relation: *horizontal* fragmentation and *vertical* fragmentation. Horizontal fragmentation splits the relation by assigning each tuple of *r* to one or more fragments. Vertical fragmentation splits the relation by decomposing the scheme *R* of relation *r*.

We shall illustrate these approaches by fragmenting the relation *account*, with the schema

Account-schema = (*account-number*, *branch-name*, *balance*)

In **horizontal fragmentation**, a relation *r* is partitioned into a number of subsets, $r_1, r_2, \ldots, r_n$. Each tuple of relation *r* must belong to at least one of the fragments, so that the original relation can be reconstructed, if needed.

As an illustration, the *account* relation can be divided into several different fragments, each of which consists of tuples of accounts belonging to a particular branch. If the banking system has only two branches—Hillside and Valleyview—then there are two different fragments:

$$account_1 = \sigma_{branch\text{-}name = \text{“Hillside”}}(account)$$
$$account_2 = \sigma_{branch\text{-}name = \text{“Valleyview”}}(account)$$

Horizontal fragmentation is usually used to keep tuples at the sites where they are used the most, to minimize data transfer.

In general, a horizontal fragment can be defined as a *selection* on the global relation *r*. That is, we use a predicate P_i to construct fragment r_i:

$$r_i = \sigma_{P_i}(r)$$

We reconstruct the relation *r* by taking the union of all fragments; that is,

$$r = r_1 \cup r_2 \cup \cdots \cup r_n$$

In our example, the fragments are disjoint. By changing the selection predicates used to construct the fragments, we can have a particular tuple of *r* appear in more than one of the r_i.

In its simplest form, vertical fragmentation is the same as decomposition (see Chapter 7). **Vertical fragmentation** of *r*(*R*) involves the definition of several subsets of attributes $R_1, R_2, \ldots, R_n$ of the schema *R* so that

$$R = R_1 \cup R_2 \cup \cdots \cup R_n$$

Each fragment r_i of *r* is defined by

$$r_i = \Pi_{R_i}(r)$$

The fragmentation should be done in such a way that we can reconstruct relation *r* from the fragments by taking the natural join

$$r = r_1 \bowtie r_2 \bowtie r_3 \bowtie \cdots \bowtie r_n$$

One way of ensuring that the relation r can be reconstructed is to include the primary-key attributes of R in each of the R_i. More generally, any superkey can be used. It is often convenient to add a special attribute, called a *tuple-id*, to the schema *R*. The tuple-id value of a tuple is a unique value that distinguishes the tuple from all other tuples. The tuple-id attribute thus serves as a candidate key for the augmented schema, and is included in each of the R_is. The physical or logical address for a tuple can be used as a tuple-id, since each tuple has a unique address.

To illustrate vertical fragmentation, consider a university database with a relation *employee-info* that stores, for each employee, *employee-id*, *name*, *designation*, and *salary*. For privacy reasons, this relation may be fragmented into a relation *employee-private-info* containing *employee-id* and *salary*, and another relation *employee-public-info* containing attributes *employee-id*, *name*, and *designation*. These may be stored at different sites, again for security reasons.

The two types of fragmentation can be applied to a single schema; for instance, the fragments obtained by horizontally fragmenting a relation can be further partitioned vertically. Fragments can also be replicated. In general, a fragment can be replicated, replicas of fragments can be fragmented further, and so on.

19.2.3 Transparency

The user of a distributed database system should not be required to know either where the data are physically located or how the data can be accessed at the specific local site. This characteristic, called **data transparency**, can take several forms:

- **Fragmentation transparency**. Users are not required to know how a relation has been fragmented.
- **Replication transparency**. Users view each data object as logically unique. The distributed system may replicate an object to increase either system per-

formance or data availability. Users do not have to be concerned with what data objects have been replicated, or where replicas have been placed.

- **Location transparency**. Users are not required to know the physical location of the data. The distributed database system should be able to find any data as long as the data identifier is supplied by the user transaction.

Data items—such as relations, fragments, and replicas — must have unique names. This property is easy to ensure in a centralized database. In a distributed database, however, we must take care to ensure that two sites do not use the same name for distinct data items.

One solution to this problem is to require all names to be registered in a central **name server**. The name server helps to ensure that the same name does not get used for different data items. We can also use the name server to locate a data item, given the name of the item. This approach, however, suffers from two major disadvantages. First, the name server may become a performance bottleneck when data items are located by their names, resulting in poor performance. Second, if the name server crashes, it may not be possible for any site in the distributed system to continue to run.

A more widely used alternative approach requires that each site prefix its own site identifier to any name that it generates. This approach ensures that no two sites generate the same name (since each site has a unique identifier). Furthermore, no central control is required. This solution, however, fails to achieve location transparency, since site identifiers are attached to names. Thus, the *account* relation might be referred to as *site*17.*account*, or *account*@*site*17, rather than as simply *account*. Many database systems use the internet address of a site to identify it.

To overcome this problem, the database system can create a set of alternative names or **aliases** for data items. A user may thus refer to data items by simple names that are translated by the system to complete names. The mapping of aliases to the real names can be stored at each site. With aliases, the user can be unaware of the physical location of a data item. Furthermore, the user will be unaffected if the database administrator decides to move a data item from one site to another.

Users should not have to refer to a specific replica of a data item. Instead, the system should determine which replica to reference on a read request, and should update all replicas on a write request. We can ensure that it does so by maintaining a catalog table, which the system uses to determine all replicas for the data item.

19.3 Distributed Transactions

Access to the various data items in a distributed system is usually accomplished through transactions, which must preserve the ACID properties (Section 15.1). There are two types of transaction that we need to consider. The **local transactions** are those that access and update data in only one local database; the **global transactions** are those that access and update data in several local databases. Ensuring the ACID properties of the local transactions can be done as described in Chapters 15, 16, and 17. However, for global transactions, this task is much more complicated, since several

sites may be participating in execution. The failure of one of these sites, or the failure of a communication link connecting these sites, may result in erroneous computations.

In this section we study the system structure of a distributed database, and its possible failure modes. On the basis of the model presented in this section, in Section 19.4 we study protocols for ensuring atomic commit of global transactions, and in Section 19.5 we study protocols for concurrency control in distributed databases. In Section 19.6 we study how a distributed database can continue functioning even in the presence of various types of failure.

19.3.1 System Structure

Each site has its own *local* transaction manager, whose function is to ensure the ACID properties of those transactions that execute at that site. The various transaction managers cooperate to execute global transactions. To understand how such a manager can be implemented, consider an abstract model of a transaction system, in which each site contains two subsystems:

- The **transaction manager** manages the execution of those transactions (or subtransactions) that access data stored in a local site. Note that each such transaction may be either a local transaction (that is, a transaction that executes at only that site) or part of a global transaction (that is, a transaction that executes at several sites).
- The **transaction coordinator** coordinates the execution of the various transactions (both local and global) initiated at that site.

The overall system architecture appears in Figure 19.1.

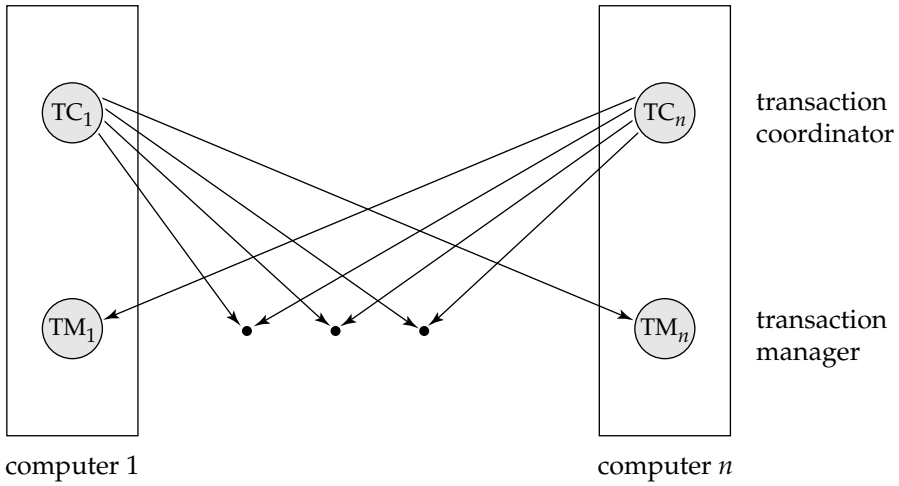

Figure 19.1 System architecture.

The structure of a transaction manager is similar in many respects to the structure of a centralized system. Each transaction manager is responsible for

- Maintaining a log for recovery purposes
- Participating in an appropriate concurrency-control scheme to coordinate the concurrent execution of the transactions executing at that site

As we shall see, we need to modify both the recovery and concurrency schemes to accommodate the distribution of transactions.

The transaction coordinator subsystem is not needed in the centralized environment, since a transaction accesses data at only a single site. A transaction coordinator, as its name implies, is responsible for coordinating the execution of all the transactions initiated at that site. For each such transaction, the coordinator is responsible for

- Starting the execution of the transaction
- Breaking the transaction into a number of subtransactions and distributing these subtransactions to the appropriate sites for execution
- Coordinating the termination of the transaction, which may result in the transaction being committed at all sites or aborted at all sites

19.3.2 System Failure Modes

A distributed system may suffer from the same types of failure that a centralized system does (for example, software errors, hardware errors, or disk crashes). There are, however, additional types of failure with which we need to deal in a distributed environment. The basic failure types are

- Failure of a site
- Loss of messages
- Failure of a communication link
- Network partition

The loss or corruption of messages is always a possibility in a distributed system. The system uses transmission-control protocols, such as TCP/IP, to handle such errors. Information about such protocols may be found in standard textbooks on networking (see the bibliographical notes).

However, if two sites A and B are not directly connected, messages from one to the other must be *routed* through a sequence of communication links. If a communication link fails, messages that would have been transmitted across the link must be rerouted. In some cases, it is possible to find another route through the network, so that the messages are able to reach their destination. In other cases, a failure may result in there being no connection between some pairs of sites. A system is **partitioned**

if it has been split into two (or more) subsystems, called **partitions**, that lack any connection between them. Note that, under this definition, a subsystem may consist of a single node.

19.4 Commit Protocols

If we are to ensure atomicity, all the sites in which a transaction T executed must agree on the final outcome of the execution. T must either commit at all sites, or it must abort at all sites. To ensure this property, the transaction coordinator of T must execute a *commit protocol*.

Among the simplest and most widely used commit protocols is the **two-phase commit protocol** (**2PC**), which is described in Section 19.4.1. An alternative is the **three-phase commit protocol** (**3PC**), which avoids certain disadvantages of the 2PC protocol but adds to complexity and overhead. Section 19.4.2 briefly outlines the 3PC protocol.

19.4.1 Two-Phase Commit

We first describe how the two-phase commit protocol (2PC) operates during normal operation, then describe how it handles failures and finally how it carries out recovery and concurrency control.

Consider a transaction T initiated at site S_i, where the transaction coordinator is C_i.

19.4.1.1 The Commit Protocol

When T completes its execution—that is, when all the sites at which T has executed inform C_i that T has completed—C_i starts the 2PC protocol.

- **Phase 1**. C_i adds the record <prepare T> to the log, and forces the log onto stable storage. It then sends a prepare T message to all sites at which T executed. On receiving such a message, the transaction manager at that site determines whether it is willing to commit its portion of T. If the answer is no, it adds a record <no T> to the log, and then responds by sending an abort T message to C_i. If the answer is yes, it adds a record <ready T> to the log, and forces the log (with all the log records corresponding to T) onto stable storage. The transaction manager then replies with a ready T message to C_i.
- **Phase 2**. When C_i receives responses to the prepare T message from all the sites, or when a prespecified interval of time has elapsed since the prepare T message was sent out, C_i can determine whether the transaction T can be committed or aborted. Transaction T can be committed if C_i received a ready T message from all the participating sites. Otherwise, transaction T must be aborted. Depending on the verdict, either a record <commit T> or a record <abort T> is added to the log and the log is forced onto stable storage. At this point, the fate of the transaction has been sealed. Following this point, the

coordinator sends either a commit T or an abort T message to all participating sites. When a site receives that message, it records the message in the log.

A site at which T executed can unconditionally abort T at any time before it sends the message ready T to the coordinator. Once the message is sent, the transaction is said to be in the **ready state** at the site. The ready T message is, in effect, a promise by a site to follow the coordinator's order to commit T or to abort T. To make such a promise, the needed information must first be stored in stable storage. Otherwise, if the site crashes after sending ready T, it may be unable to make good on its promise. Further, locks acquired by the transaction must continue to be held till the transaction completes.

Since unanimity is required to commit a transaction, the fate of T is sealed as soon as at least one site responds abort T. Since the coordinator site S_i is one of the sites at which T executed, the coordinator can decide unilaterally to abort T. The final verdict regarding T is determined at the time that the coordinator writes that verdict (commit or abort) to the log and forces that verdict to stable storage. In some implementations of the 2PC protocol, a site sends an acknowledge T message to the coordinator at the end of the second phase of the protocol. When the coordinator receives the acknowledge T message from all the sites, it adds the record <complete T> to the log.

19.4.1.2 Handling of Failures

The 2PC protocol responds in differenct ways to various types of failures:

- **Failure of a participating site**. If the coordinator C_i detects that a site has failed, it takes these actions: If the site fails before responding with a ready T message to C_i, the coordinator assumes that it responded with an abort T message. If the site fails after the coordinator has received the ready T message from the site, the coordinator executes the rest of the commit protocol in the normal fashion, ignoring the failure of the site.

  When a participating site S_k recovers from a failure, it must examine its log to determine the fate of those transactions that were in the midst of execution when the failure occurred. Let T be one such transaction. We consider each of the possible cases:
  - ☐ The log contains a <commit T> record. In this case, the site executes redo(T).
  - ☐ The log contains an <abort T> record. In this case, the site executes undo(T).
  - ☐ The log contains a <ready T> record. In this case, the site must consult C_i to determine the fate of T. If C_i is up, it notifies S_k regarding whether T committed or aborted. In the former case, it executes redo(T); in the latter case, it executes undo(T). If C_i is down, S_k must try to find the fate of T from other sites. It does so by sending a querystatus T message to all the sites in the system. On receiving such a message, a site must consult its log to determine whether T has executed there, and if T has, whether T committed or aborted. It then notifies S_k about this outcome. If no site has the appropriate information (that is, whether T committed or aborted), then S_k can neither abort nor commit T. The decision concerning T is

postponed until S_k can obtain the needed information. Thus, S_k must periodically resend the querystatus message to the other sites. It continues to do so until a site that contains the needed information recovers. Note that the site at which C_i resides always has the needed information.

- The log contains no control records (abort, commit, ready) concerning T. Thus, we know that S_k failed before responding to the prepare T message from C_i. Since the failure of S_k precludes the sending of such a response, by our algorithm C_i must abort T. Hence, S_k must execute undo(T).

- **Failure of the coordinator**. If the coordinator fails in the midst of the execution of the commit protocol for transaction T, then the participating sites must decide the fate of T. We shall see that, in certain cases, the participating sites cannot decide whether to commit or abort T, and therefore these sites must wait for the recovery of the failed coordinator.
  - If an active site contains a <commit T> record in its log, then T must be committed.
  - If an active site contains an <abort T> record in its log, then T must be aborted.
  - If some active site does *not* contain a <ready T> record in its log, then the failed coordinator C_i cannot have decided to commit T, because a site that does not have a <ready T> record in its log cannot have sent a ready T message to C_i. However, the coordinator may have decided to abort T, but not to commit T. Rather than wait for C_i to recover, it is preferable to abort T.
  - If none of the preceding cases holds, then all active sites must have a <ready T> record in their logs, but no additional control records (such as <abort T> or <commit T>). Since the coordinator has failed, it is impossible to determine whether a decision has been made, and if one has, what that decision is, until the coordinator recovers. Thus, the active sites must wait for C_i to recover. Since the fate of T remains in doubt, T may continue to hold system resources. For example, if locking is used, T may hold locks on data at active sites. Such a situation is undesirable, because it may be hours or days before C_i is again active. During this time, other transactions may be forced to wait for T. As a result, data items may be unavailable not only on the failed site (C_i), but on active sites as well. This situation is called the **blocking** problem, because T is blocked pending the recovery of site C_i.

- **Network partition**. When a network partitions, two possibilities exist:
  1. The coordinator and all its participants remain in one partition. In this case, the failure has no effect on the commit protocol.
  2. The coordinator and its participants belong to several partitions. From the viewpoint of the sites in one of the partitions, it appears that the sites in other partitions have failed. Sites that are not in the partition containing the coordinator simply execute the protocol to deal with failure of the coordinator. The coordinator and the sites that are in the same partition as

the coordinator follow the usual commit protocol, assuming that the sites in the other partitions have failed.

Thus, the major disadvantage of the 2PC protocol is that coordinator failure may result in blocking, where a decision either to commit or to abort T may have to be postponed until C_i recovers.

19.4.1.3 Recovery and Concurrency Control

When a failed site restarts, we can perform recovery by using, for example, the recovery algorithm described in Section 17.9. To deal with distributed commit protocols (such as 2PC and 3PC), the recovery procedure must treat **in-doubt transactions** specially; in-doubt transactions are transactions for which a <ready T> log record is found, but neither a <commit T> log record nor an <abort T> log record is found. The recovering site must determine the commit–abort status of such transactions by contacting other sites, as described in Section 19.4.1.2.

If recovery is done as just described, however, normal transaction processing at the site cannot begin until all in-doubt transactions have been committed or rolled back. Finding the status of in-doubt transactions can be slow, since multiple sites may have to be contacted. Further, if the coordinator has failed, and no other site has information about the commit–abort status of an incomplete transaction, recovery potentially could become blocked if 2PC is used. As a result, the site performing restart recovery may remain unusable for a long period.

To circumvent this problem, recovery algorithms typically provide support for noting lock information in the log. (We are assuming here that locking is used for concurrency control.) Instead of writing a <ready T> log record, the algorithm writes a <ready T, L> log record, where L is a list of all write locks held by the transaction T when the log record is written. At recovery time, after performing local recovery actions, for every in-doubt transaction T, all the write locks noted in the <ready T, L> log record (read from the log) are reacquired.

After lock reacquisition is complete for all in-doubt transactions, transaction processing can start at the site, even before the commit–abort status of the in-doubt transactions is determined. The commit or rollback of in-doubt transactions proceeds concurrently with the execution of new transactions. Thus, site recovery is faster, and never gets blocked. Note that new transactions that have a lock conflict with any write locks held by in-doubt transactions will be unable to make progress until the conflicting in-doubt transactions have been committed or rolled back.

19.4.2 Three-Phase Commit

The three-phase commit (3PC) protocol is an extension of the two-phase commit protocol that avoids the blocking problem under certain assumptions. In particular, it is assumed that no network partition occurs, and not more than k sites fail, where k is some predetermined number. Under these assumptions, the protocol avoids blocking by introducing an extra third phase where multiple sites are involved in the decision to commit. Instead of directly noting the commit decision in its persistent storage, the

coordinator first ensures that at least k other sites know that it intended to commit the transaction. If the coordinator fails, the remaining sites first select a new coordinator. This new coordinator checks the status of the protocol from the remaining sites; if the coordinator had decided to commit, at least one of the other k sites that it informed will be up and will ensure that the commit decision is respected. The new coordinator restarts the third phase of the protocol if some site knew that the old coordinator intended to commit the transaction. Otherwise the new coordinator aborts the transaction.

While the 3PC protocol has the desirable property of not blocking unless k sites fail, it has the drawback that a partitioning of the network will appear to be the same as more than k sites failing, which would lead to blocking. The protocol also has to be carefully implemented to ensure that network partitioning (or more than k sites failing) does not result in inconsistencies, where a transaction is committed in one partition, and aborted in another. Because of its overhead, the 3PC protocol is not widely used. See the bibliographical notes for references giving more details of the 3PC protocol.

19.4.3 Alternative Models of Transaction Processing

For many applications, the blocking problem of two-phase commit is not acceptable. The problem here is the notion of a single transaction that works across multiple sites. In this section we describe how to use *persistent messaging* to avoid the problem of distributed commit, and then briefly outline the larger issue of *workflows*; workflows are considered in more detail in Section 24.2.

To understand persistent messaging consider how one might transfer funds between two different banks, each with its own computer. One approach is to have a transaction span the two sites, and use two-phase commit to ensure atomicity. However, the transaction may have to update the total bank balance, and blocking could have a serious impact on all other transactions at each bank, since almost all transactions at the bank would update the total bank balance.

In contrast, consider how fund transfer by a bank check occurs. The bank first deducts the amount of the check from the available balance and prints out a check. The check is then physically transferred to the other bank where it is deposited. After verifying the check, the bank increases the local balance by the amount of the check. The check constitutes a message sent between the two banks. So that funds are not lost or incorrectly increased, the check must not be lost, and must not be duplicated and deposited more than once. When the bank computers are connected by a network, persistent messages provide the same service as the check (but much faster, of course).

Persistent messages are messages that are guaranteed to be delivered to the recipient exactly once (neither less nor more), regardless of failures, if the transaction sending the message commits, and are guaranteed to not be delivered if the transaction aborts. Database recovery techniques are used to implement persistent messaging on top of the normal network channels, as we will see shortly. In contrast, regular messages may be lost or may even be delivered multiple times in some situations.

Error handling is more complicated with persistent messaging than with two-phase commit. For instance, if the account where the check is to be deposited has been closed, the check must be sent back to the originating account and credited back there. Both sites must therefore be provided with error handling code, along with code to handle the persistent messages. In contrast, with two-phase commit, the error would be detected by the transaction, which would then never deduct the amount in the first place.

The types of exception conditions that may arise depend on the application, so it is not possible for the database system to handle exceptions automatically. The application programs that send and receive persistent messages must include code to handle exception conditions and bring the system back to a consistent state. For instance, it is not acceptable to just lose the money being transfered if the receiving account has been closed; the money must be credited back to the originating account, and if that is not possible for some reason, humans must be alerted to resolve the situation manually.

There are many applications where the benefit of eliminating blocking is well worth the extra effort to implement systems that use persistent messages. In fact, few organizations would agree to support two-phase commit for transactions originating outside the organization, since failures could result in blocking of access to local data. Persistent messaging therefore plays an important role in carrying out transactions that cross organizational boundaries.

Workflows provide a general model of transaction processing involving multiple sites and possibly human processing of certain steps. For instance, when a bank receives a loan application, there are many steps it must take, including contacting external credit-checking agencies, before approving or rejecting a loan application. The steps, together, form a workflow. We study workflows in more detail in Section 24.2. We also note that persistent messaging forms the underlying basis for workflows in a distributed environment.

We now consider the **implementation** of persistent messaging. Persistent messaging can be implemented on top of an unreliable messaging infrastructure, which may lose messages or deliver them multiple times, by these protocols:

- **Sending site protocol**: When a transaction wishes to send a persistent message, it writes a record containing the message in a special relation *messages-to-send*, instead of directly sending out the message. The message is also given a unique message identifier.

  A *message delivery process* monitors the relation, and when a new message is found, it sends the message to its destination. The usual database concurrency control mechanisms ensure that the system process reads the message only after the transaction that wrote the message commits; if the transaction aborts, the usual recovery mechanism would delete the message from the relation.

  The message delivery process deletes a message from the relation only after it receives an acknowledgment from the destination site. If it receives no acknowledgement from the destination site, after some time it sends the message again. It repeats this until an acknowledgment is received. In case of permanent failures, the system will decide, after some period of time, that the

message is undeliverable. Exception handling code provided by the application is then invoked to deal with the failure.

Writing the message to a relation and processing it only after the transaction commits ensures that the message will be delivered if and only if the transaction commits. Repeatedly sending it guarantees it will be delivered even if there are (temporary) system or network failures.

- **Receiving site protocol**: When a site receives a persistent message, it runs a transaction that adds the message to a special *received-messages* relation, provided it is not already present in the relation (the unique message identifier detects duplicates). After the transaction commits, or if the message was already present in the relation, the receiving site sends an acknowledgment back to the sending site.

  Note that sending the acknowledgment before the transaction commits is not safe, since a system failure may then result in loss of the message. Checking whether the message has been received earlier is essential to avoid multiple deliveries of the message.

  In many messaging systems, it is possible for messages to get delayed arbitrarily, although such delays are very unlikely. Therefore, to be safe, the message must never be deleted from the *received-messages* relation. Deleting it could result in a duplicate delivery not being detected. But as a result, the *received-messages* relation may grow indefinitely. To deal with this problem, each message is given a timestamp, and if the timestamp of a received message is older than some cutoff, the message is discarded. All messages recorded in the *received-messages* relation that are older than the cutoff can be deleted.

19.5 Concurrency Control in Distributed Databases

We show here how some of the concurrency-control schemes discussed in Chapter 16 can be modified so that they can be used in a distributed environment. We assume that each site participates in the execution of a commit protocol to ensure global transaction atomicity.

The protocols we describe in this section require updates to be done on all replicas of a data item. If any site containing a replica of a data item has failed, updates to the data item cannot be processed. In Section 19.6 we describe protocols that can continue transaction processing even if some sites or links have failed, thereby providing high availability.

19.5.1 Locking Protocols

The various locking protocols described in Chapter 16 can be used in a distributed environment. The only change that needs to be incorporated is in the way the lock manager deals with replicated data. We present several possible schemes that are applicable to an environment where data can be replicated in several sites. As in Chapter 16, we shall assume the existence of the *shared* and *exclusive* lock modes.

19.5.1.1 Single Lock-Manager Approach

In the **single lock-manager** approach, the system maintains a *single* lock manager that resides in a *single* chosen site—say S_i. All lock and unlock requests are made at site S_i. When a transaction needs to lock a data item, it sends a lock request to S_i. The lock manager determines whether the lock can be granted immediately. If the lock can be granted, the lock manager sends a message to that effect to the site at which the lock request was initiated. Otherwise, the request is delayed until it can be granted, at which time a message is sent to the site at which the lock request was initiated. The transaction can read the data item from *any* one of the sites at which a replica of the data item resides. In the case of a write, all the sites where a replica of the data item resides must be involved in the writing.

The scheme has these advantages:

- **Simple implementation**. This scheme requires two messages for handling lock requests, and one message for handling unlock requests.
- **Simple deadlock handling**. Since all lock and unlock requests are made at one site, the deadlock-handling algorithms discussed in Chapter 16 can be applied directly to this environment.

The disadvantages of the scheme are:

- **Bottleneck**. The site S_i becomes a bottleneck, since all requests must be processed there.
- **Vulnerability**. If the site S_i fails, the concurrency controller is lost. Either processing must stop, or a recovery scheme must be used so that a backup site can take over lock management from S_i, as described in Section 19.6.5.

19.5.1.2 Distributed Lock Manager

A compromise between the advantages and disadvantages can be achieved through the **distributed lock-manager** approach, in which the lock-manager function is distributed over several sites.

Each site maintains a local lock manager whose function is to administer the lock and unlock requests for those data items that are stored in that site. When a transaction wishes to lock data item Q, which is not replicated and resides at site S_i, a message is sent to the lock manager at site S_i requesting a lock (in a particular lock mode). If data item Q is locked in an incompatible mode, then the request is delayed until it can be granted. Once it has determined that the lock request can be granted, the lock manager sends a message back to the initiator indicating that it has granted the lock request.

There are several alternative ways of dealing with replication of data items, which we study in Sections 19.5.1.3 to 19.5.1.6.

The distributed lock manager scheme has the advantage of simple implementation, and reduces the degree to which the coordinator is a bottleneck. It has a reasonably low overhead, requiring two message transfers for handling lock requests, and

one message transfer for handling unlock requests. However, deadlock handling is more complex, since the lock and unlock requests are no longer made at a single site: There may be intersite deadlocks even when there is no deadlock within a single site. The deadlock-handling algorithms discussed in Chapter 16 must be modified, as we shall discuss in Section 19.5.4, to detect global deadlocks.

19.5.1.3 Primary Copy

When a system uses data replication, we can choose one of the replicas as the **primary copy**. Thus, for each data item Q, the primary copy of Q must reside in precisely one site, which we call the **primary site** of Q.

When a transaction needs to lock a data item Q, it requests a lock at the primary site of Q. As before, the response to the request is delayed until it can be granted.

Thus, the primary copy enables concurrency control for replicated data to be handled like that for unreplicated data. This similarity allows for a simple implementation. However, if the primary site of Q fails, Q is inaccessible, even though other sites containing a replica may be accessible.

19.5.1.4 Majority Protocol

The **majority protocol** works this way: If data item Q is replicated in n different sites, then a lock-request message must be sent to more than one-half of the n sites in which Q is stored. Each lock manager determines whether the lock can be granted immediately (as far as it is concerned). As before, the response is delayed until the request can be granted. The transaction does not operate on Q until it has successfully obtained a lock on a majority of the replicas of Q.

This scheme deals with replicated data in a decentralized manner, thus avoiding the drawbacks of central control. However, it suffers from these disadvantages:

- **Implementation**. The majority protocol is more complicated to implement than are the previous schemes. It requires $2(n/2 + 1)$ messages for handling lock requests, and $(n/2 + 1)$ messages for handling unlock requests.
- **Deadlock handling**. In addition to the problem of global deadlocks due to the use of a distributed lock-manager approach, it is possible for a deadlock to occur even if only one data item is being locked. As an illustration, consider a system with four sites and full replication. Suppose that transactions T_1 and T_2 wish to lock data item Q in exclusive mode. Transaction T_1 may succeed in locking Q at sites S_1 and S_3, while transaction T_2 may succeed in locking Q at sites S_2 and S_4. Each then must wait to acquire the third lock; hence, a deadlock has occurred. Luckily, we can avoid such deadlocks with relative ease, by requiring all sites to request locks on the replicas of a data item in the same predetermined order.

19.5.1.5 Biased Protocol

The **biased protocol** is another approach to handling replication. The difference from the majority protocol is that requests for shared locks are given more favorable treatment than requests for exclusive locks.

- **Shared locks**. When a transaction needs to lock data item Q, it simply requests a lock on Q from the lock manager at one site that contains a replica of Q.
- **Exclusive locks**. When a transaction needs to lock data item Q, it requests a lock on Q from the lock manager at all sites that contain a replica of Q.

As before, the response to the request is delayed until it can be granted.

The biased scheme has the advantage of imposing less overhead on read operations than does the majority protocol. This savings is especially significant in common cases in which the frequency of read is much greater than the frequency of write. However, the additional overhead on writes is a disadvantage. Furthermore, the biased protocol shares the majority protocol's disadvantage of complexity in handling deadlock.

19.5.1.6 Quorum Consensus Protocol

The **quorum consensus** protocol is a generalization of the majority protocol. The quorum consensus protocol assigns each site a nonnegative weight. It assigns read and write operations on an item x two integers, called **read quorum** Q_r and **write quorum** Q_w, that must satisfy the following condition, where S is the total weight of all sites at which x resides:

$$Q_r + Q_w > S \textbf{ and } 2 * Q_w > S$$

To execute a read operation, enough replicas must be read that their total weight is $\geq Q_r$. To execute a write operation, enough replicas must be written so that their total weight is $\geq Q_w$.

The benefit of the quorum consensus approach is that it can permit the cost of either reads or writes to be selectively reduced by appropriately defining the read and write quorums. For instance, with a small read quorum, reads need to read fewer replicas, but the write quorum will be higher, hence writes can succeed only if correspondingly more replicas are available. Also, if higher weights are given to some sites (for example, those less likely to fail), fewer sites need to be accessed for acquiring locks.

In fact, by setting weights and quorums appropriately, the quorum consensus protocol can simulate the majority protocol and the biased protocols.

19.5.2 Timestamping

The principal idea behind the timestamping scheme in Section 16.2 is that each transaction is given a *unique* timestamp that the system uses in deciding the serialization order. Our first task, then, in generalizing the centralized scheme to a distributed

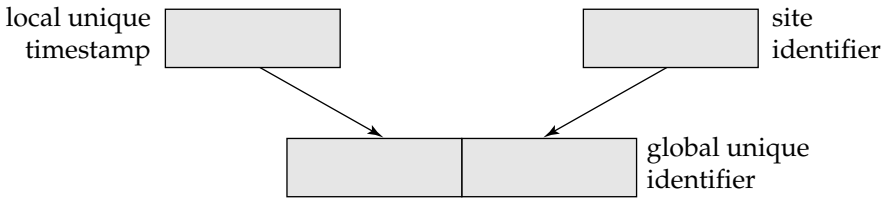

Figure 19.2 Generation of unique timestamps.

scheme is to develop a scheme for generating unique timestamps. Then, the various protocols can operate directly to the nonreplicated environment.

There are two primary methods for generating unique timestamps, one centralized and one distributed. In the centralized scheme, a single site distributes the timestamps. The site can use a logical counter or its own local clock for this purpose.

In the distributed scheme, each site generates a unique local timestamp by using either a logical counter or the local clock. We obtain the unique global timestamp by concatenating the unique local timestamp with the site identifier, which also must be unique (Figure 19.2). The order of concatenation is important! We use the site identifier in the least significant position to ensure that the global timestamps generated in one site are not always greater than those generated in another site. Compare this technique for generating unique timestamps with the one that we presented in Section 19.2.3 for generating unique names.

We may still have a problem if one site generates local timestamps at a rate faster than that of the other sites. In such a case, the fast site's logical counter will be larger than that of other sites. Therefore, all timestamps generated by the fast site will be larger than those generated by other sites. What we need is a mechanism to ensure that local timestamps are generated fairly across the system. We define within each site S_i a **logical clock** (LC_i), which generates the unique local timestamp. The logical clock can be implemented as a counter that is incremented after a new local timestamp is generated. To ensure that the various logical clocks are synchronized, we require that a site S_i advance its logical clock whenever a transaction T_i with timestamp $<x,y>$ visits that site and x is greater than the current value of LC_i. In this case, site S_i advances its logical clock to the value $x + 1$.

If the system clock is used to generate timestamps, then timestamps will be assigned fairly, provided that no site has a system clock that runs fast or slow. Since clocks may not be perfectly accurate, a technique similar to that for logical clocks must be used to ensure that no clock gets far ahead of or behind another clock.

19.5.3 Replication with Weak Degrees of Consistency

Many commercial databases today support replication, which can take one of several forms. With **master–slave replication**, the database allows updates at a primary site, and automatically propagates updates to replicas at other sites. Transactions may read the replicas at other sites, but are not permitted to update them.

An important feature of such replication is that transactions do not obtain locks at remote sites. To ensure that transactions running at the replica sites see a consistent

(but perhaps outdated) view of the database, the replica should reflect a **transaction-consistent snapshot** of the data at the primary; that is, the replica should reflect all updates of transactions up to some transaction in the serialization order, and should not reflect any updates of later transactions in the serialization order.

The database may be configured to propagate updates immediately after they occur at the primary, or to propagate updates only periodically.

Master–slave replication is particularly useful for distributing information, for instance from a central office to branch offices of an organization. Another use for this form of replication is in creating a copy of the database to run large queries, so that queries do not interfere with transactions. Updates should be propagated periodically—every night, for example—so that update propagation does not interfere with query processing.

The Oracle database system supports a **create snapshot** statement, which can create a transaction-consistent snapshot copy of a relation, or set of relations, at a remote site. It also supports snapshot refresh, which can be done either by recomputing the snapshot or by incrementally updating it. Oracle supports automatic refresh, either continuously or at periodic intervals.

With **multimaster replication** (also called **update-anywhere replication**) updates are permitted at any replica of a data item, and are automatically propagated to all replicas. This model is the basic model used to manage replicas in distributed databases. Transactions update the local copy and the system updates other replicas transparently.

One way of updating replicas is to apply immediate update with two-phase commit, using one of the distributed concurrency-control techniques we have seen. Many database systems use the biased protocol, where writes have to lock and update all replicas and reads lock and read any one replica, as their currency-control technique.

Many database systems provide an alternative form of updating: They update at one site, with **lazy propagation** of updates to other sites, instead of immediately applying updates to all replicas as part of the transaction performing the update. Schemes based on lazy propagation allow transaction processing (including updates) to proceed even if a site is disconnected from the network, thus improving availability, but, unfortunately, do so at the cost of consistency. One of two approaches is usually followed when lazy propagation is used:

- Updates at replicas are translated into updates at a primary site, which are then propagated lazily to all replicas.

  This approach ensures that updates to an item are ordered serially, although serializability problems can occur, since transactions may read an old value of some other data item and use it to perform an update.

- Updates are performed at any replica and propagated to all other replicas.

  This approach can cause even more problems, since the same data item may be updated concurrently at multiple sites.

Some conflicts due to the lack of distributed concurrency control can be detected when updates are propagated to other sites (we shall see how in Section 23.5.4),

but resolving the conflict involves rolling back committed transactions, and durability of committed transactions is therefore not guaranteed. Further, human intervention may be required to deal with conflicts. The above schemes should therefore be avoided or used with care.

19.5.4 Deadlock Handling

The deadlock-prevention and deadlock-detection algorithms in Chapter 16 can be used in a distributed system, provided that modifications are made. For example, we can use the tree protocol by defining a *global* tree among the system data items. Similarly, the timestamp-ordering approach could be directly applied to a distributed environment, as we saw in Section 19.5.2.

Deadlock prevention may result in unnecessary waiting and rollback. Furthermore, certain deadlock-prevention techniques may require more sites to be involved in the execution of a transaction than would otherwise be the case.

If we allow deadlocks to occur and rely on deadlock detection, the main problem in a distributed system is deciding how to maintain the wait-for graph. Common techniques for dealing with this issue require that each site keep a **local wait-for graph**. The nodes of the graph correspond to all the transactions (local as well as nonlocal) that are currently either holding or requesting any of the items local to that site. For example, Figure 19.3 depicts a system consisting of two sites, each maintaining its local wait-for graph. Note that transactions T_2 and T_3 appear in both graphs, indicating that the transactions have requested items at both sites.

These local wait-for graphs are constructed in the usual manner for local transactions and data items. When a transaction T_i on site S_1 needs a resource in site S_2, it sends a request message to site S_2. If the resource is held by transaction T_j, the system inserts an edge $T_i \rightarrow T_j$ in the local wait-for graph of site S_2.

Clearly, if any local wait-for graph has a cycle, deadlock has occurred. On the other hand, the fact that there are no cycles in any of the local wait-for graphs does not mean that there are no deadlocks. To illustrate this problem, we consider the local wait-for graphs of Figure 19.3. Each wait-for graph is acyclic; nevertheless, a deadlock exists in the system because the *union* of the local wait-for graphs contains a cycle. This graph appears in Figure 19.4.

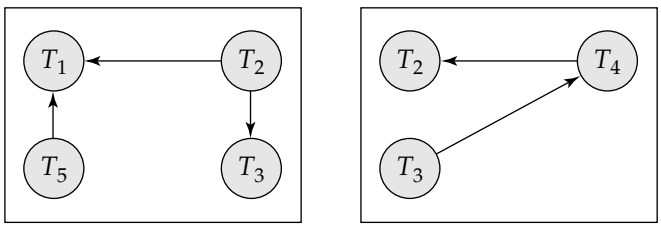

Figure 19.3 Local wait-for graphs.

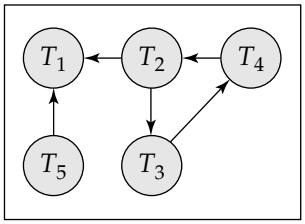

Figure 19.4 Global wait-for graph for Figure 19.3.

In the **centralized deadlock detection** approach, the system constructs and maintains a **global wait-for graph** (the union of all the local graphs) in a *single* site: the deadlock-detection coordinator. Since there is communication delay in the system, we must distinguish between two types of wait-for graphs. The *real* graph describes the real but unknown state of the system at any instance in time, as would be seen by an omniscient observer. The *constructed* graph is an approximation generated by the controller during the execution of the controller's algorithm. Obviously, the controller must generate the constructed graph in such a way that, whenever the detection algorithm is invoked, the reported results are correct. *Correct* means in this case that, if a deadlock exists, it is reported promptly, and if the system reports a deadlock, it is indeed in a deadlock state.

The global wait-for graph can be reconstructed or updated under these conditions:

- Whenever a new edge is inserted in or removed from one of the local wait-for graphs.
- Periodically, when a number of changes have occurred in a local wait-for graph.
- Whenever the coordinator needs to invoke the cycle-detection algorithm.

When the coordinator invokes the deadlock-detection algorithm, it searches its global graph. If it finds a cycle, it selects a victim to be rolled back. The coordinator must notify all the sites that a particular transaction has been selected as victim. The sites, in turn, roll back the victim transaction.

This scheme may produce unnecessary rollbacks if:

- **False cycles** exist in the global wait-for graph. As an illustration, consider a snapshot of the system represented by the local wait-for graphs of Figure 19.5. Suppose that T_2 releases the resource that it is holding in site S_1, resulting in the deletion of the edge $T_1 \rightarrow T_2$ in S_1. Transaction T_2 then requests a resource held by T_3 at site S_2, resulting in the addition of the edge $T_2 \rightarrow T_3$ in S_2. If the insert $T_2 \rightarrow T_3$ message from S_2 arrives before the remove $T_1 \rightarrow T_2$ message from S_1, the coordinator may discover the false cycle $T_1 \rightarrow T_2 \rightarrow T_3$ after the insert (but before the remove). Deadlock recovery may be initiated, although no deadlock has occurred.

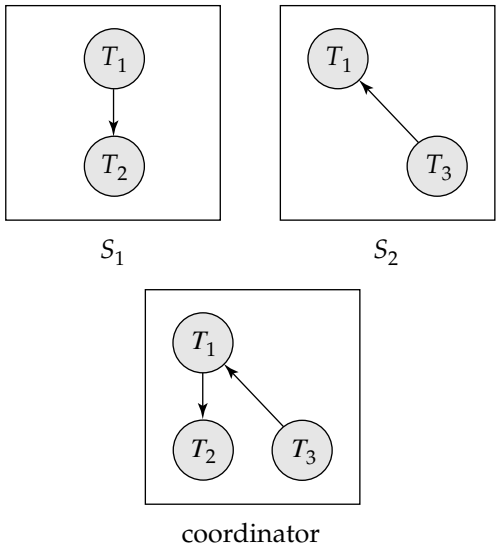

Figure 19.5 False cycles in the global wait-for graph.

Note that the false-cycle situation could not occur under two-phase locking. The likelihood of false cycles is usually sufficiently low that they do not cause a serious performance problem.

- A *deadlock* has indeed occurred and a victim has been picked, while one of the transactions was aborted for reasons unrelated to the deadlock. For example, suppose that site S_1 in Figure 19.3 decides to abort T_2. At the same time, the coordinator has discovered a cycle, and has picked T_3 as a victim. Both T_2 and T_3 are now rolled back, although only T_2 needed to be rolled back.

Deadlock detection can be done in a distributed manner, with several sites taking on parts of the task, instead of being done at a single site, However, such algorithms are more complicated and more expensive. See the bibliographical notes for references to such algorithms.

19.6 Availability

One of the goals in using distributed databases is **high availability**; that is, the database must function almost all the time. In particular, since failures are more likely in large distributed systems, a distributed database must continue functioning even when there are various types of failures. The ability to continue functioning even during failures is referred to as **robustness**.

For a distributed system to be robust, it must *detect* failures, *reconfigure* the system so that computation may continue, and *recover* when a processor or a link is repaired.

The different types of failures are handled in different ways. For example, message loss is handled by retransmission. Repeated retransmission of a message across a link,

without receipt of an acknowledgment, is usually a symptom of a link failure. The network usually attempts to find an alternative route for the message. Failure to find such a route is usually a symptom of network partition.

It is generally not possible, however, to differentiate clearly between site failure and network partition. The system can usually detect that a failure has occurred, but it may not be able to identify the type of failure. For example, suppose that site S_1 is not able to communicate with S_2. It could be that S_2 has failed. However, another possibility is that the link between S_1 and S_2 has failed, resulting in network partition. The problem is partly addressed by using multiple links between sites, so that even if one link fails the sites will remain connected. However, multiple link failure can still occur, so there are situations where we cannot be sure whether a site failure or network partition has occurred.

Suppose that site S_1 has discovered that a failure has occurred. It must then initiate a procedure that will allow the system to reconfigure, and to continue with the normal mode of operation.

- If transactions were active at a failed/inaccessible site at the time of the failure, these transactions should be aborted. It is desirable to abort such transactions promptly, since they may hold locks on data at sites that are still active; waiting for the failed/inaccessible site to become accessible again may impede other transactions at sites that are operational.

  However, in some cases, when data objects are replicated it may be possible to proceed with reads and updates even though some replicas are inaccessible. In this case, when a failed site recovers, if it had replicas of any data object, it must obtain the current values of these data objects, and must ensure that it receives all future updates. We address this issue in Section 19.6.1.

- If replicated data are stored at a failed/inaccessible site, the catalog should be updated so that queries do not reference the copy at the failed site. When a site rejoins, care must be taken to ensure that data at the site is consistent, as we will see in Section 19.6.3.

- If a failed site is a central server for some subsystem, an *election* must be held to determine the new server (see Section 19.6.5). Examples of central servers include a name server, a concurrency coordinator, or a global deadlock detector.

Since it is, in general, not possible to distinguish between network link failures and site failures, any reconfiguration scheme must be designed to work correctly in case of a partitioning of the network. In particular, these situations must be avoided:

- Two or more central servers are elected in distinct partitions.

- More than one partition updates a replicated data item.

19.6.1 Majority-Based Approach

The majority-based approach to distributed concurrency control in Section 19.5.1.4 can be modified to work in spite of failures. In this approach, each data object stores with it a version number to detect when it was last written to. Whenever a transaction writes an object it also updates the version number in this way:

- If data object a is replicated in n different sites, then a lock-request message must be sent to more than one-half of the n sites in which a is stored. The transaction does not operate on a until it has successfully obtained a lock on a majority of the replicas of a.
- Read operations look at all replicas on which a lock has been obtained, and read the value from the replica that has the highest version number. (Optionally, they may also write this value back to replicas with lower version numbers.) Writes read all the replicas just like reads to find the highest version number (this step would normally have been performed earlier in the transaction by a read, and the result can be reused). The new version number is one more than the highest version number. The write operation writes all the replicas on which it has obtained locks, and sets the version number at all the replicas to the new version number.

Failures during a transaction (whether network partitions or site failures) can be tolerated as long as (1) the sites available at commit contain a majority of replicas of all the objects written to and (2) during reads, a majority of replicas are read to find the version numbers. If these requirements are violated, the transaction must be aborted. As long as the requirements are satisfied, the two-phase commit protocol can be used, as usual, on the sites that are available.

In this scheme, reintegration is trivial; nothing needs to be done. This is because writes would have updated a majority of the replicas, while reads will read a majority of the replicas and find at least one replica that has the latest version.

The version numbering technique used with the majority protocol can also be used to make the quorum consensus protocol work in the presence of failures. We leave the (straightforward) details to the reader. However, the danger of failures preventing the system from processing transactions increases if some sites are given higher weights.

19.6.2 Read One, Write All Available Approach

As a special case of quorum consensus, we can employ the biased protocol by giving unit weights to all sites, setting the read quorum to 1, and setting the write quorum to n (all sites). In this special case, there is no need to use version numbers; however, if even a single site containing a data item fails, no write to the item can proceed, since the write quorum will not be available. This protocol is called the **read one, write all** protocol since all replicas must be written.

To allow work to proceed in the event of failures, we would like to be able to use a **read one, write all available** protocol. In this approach, a read operation proceeds as in the **read one, write all** scheme; any available replica can be read, and a read lock is

obtained at that replica. A write operation is shipped to all replicas; and write locks are acquired on all the replicas. If a site is down, the transaction manager proceeds without waiting for the site to recover.

While this approach appears very attractive, there are several complications. In particular, temporary communication failure may cause a site to appear to be unavailable, resulting in a write not being performed, but when the link is restored, the site is not aware that it has to perform some reintegration actions to catch up on writes it has lost. Further, if the network partitions, each partition may proceed to update the same data item, believing that sites in the other partitions are all dead.

The read one, write all available scheme can be used if there is never any network partitioning, but it can result in inconsistencies in the event of network partitions.

19.6.3 Site Reintegration

Reintegration of a repaired site or link into the system requires care. When a failed site recovers, it must initiate a procedure to update its system tables to reflect changes made while it was down. If the site had replicas of any data items, it must obtain the current values of these data items and ensure that it receives all future updates. Reintegration of a site is more complicated than it may seem to be at first glance, since there may be updates to the data items processed during the time that the site is recovering.

An easy solution is to halt the entire system temporarily while the failed site rejoins it. In most applications, however, such a temporary halt is unacceptably disruptive. Techniques have been developed to allow failed sites to reintegrate while concurrent updates to data items proceed concurrently. Before a read or write lock is granted on any data item, the site must ensure that it has caught up on all updates to the data item. If a failed link recovers, two or more partitions can be rejoined. Since a partitioning of the network limits the allowable operations by some or all sites, all sites should be informed promptly of the recovery of the link. See the bibliographical notes for more information on recovery in distributed systems.

19.6.4 Comparison with Remote Backup

Remote backup systems, which we studied in Section 17.10, and replication in distributed databases are two alternative approaches to providing high availability. The main difference between the two schemes is that with remote backup systems, actions such as concurrency control and recovery are performed at a single site, and only data and log records are replicated at the other site. In particular, remote backup systems help avoid two-phase commit, and its resultant overheads. Also, transactions need to contact only one site (the primary site), and thus avoid the overhead of running transaction code at multiple sites. Thus remote backup systems offer a lower-cost approach to high availability than replication.

On the other hand, replication can provide greater availability by having multiple replicas available, and using the majority protocol.

19.6.5 Coordinator Selection

Several of the algorithms that we have presented require the use of a coordinator. If the coordinator fails because of a failure of the site at which it resides, the system can continue execution only by restarting a new coordinator on another site. One way to continue execution is by maintaining a backup to the coordinator, which is ready to assume responsibility if the coordinator fails.

A **backup coordinator** is a site that, in addition to other tasks, maintains enough information locally to allow it to assume the role of coordinator with minimal disruption to the distributed system. All messages directed to the coordinator are received by both the coordinator and its backup. The backup coordinator executes the same algorithms and maintains the same internal state information (such as, for a concurrency coordinator, the lock table) as does the actual coordinator. The only difference in function between the coordinator and its backup is that the backup does not take any action that affects other sites. Such actions are left to the actual coordinator.

In the event that the backup coordinator detects the failure of the actual coordinator, it assumes the role of coordinator. Since the backup has all the information available to it that the failed coordinator had, processing can continue without interruption.

The prime advantage to the backup approach is the ability to continue processing immediately. If a backup were not ready to assume the coordinator's responsibility, a newly appointed coordinator would have to seek information from all sites in the system so that it could execute the coordination tasks. Frequently, the only source of some of the requisite information is the failed coordinator. In this case, it may be necessary to abort several (or all) active transactions, and to restart them under the control of the new coordinator.

Thus, the backup-coordinator approach avoids a substantial amount of delay while the distributed system recovers from a coordinator failure. The disadvantage is the overhead of duplicate execution of the coordinator's tasks. Furthermore, a coordinator and its backup need to communicate regularly to ensure that their activities are synchronized.

In short, the backup-coordinator approach incurs overhead during normal processing to allow fast recovery from a coordinator failure.

In the absence of a designated backup coordinator, or in order to handle multiple failures, a new coordinator may be chosen dynamically by sites that are live. **Election algorithms** enable the sites to choose the site for the new coordinator in a decentralized manner. Election algorithms require that a unique identification number be associated with each active site in the system.

The **bully algorithm** for election works as follows. To keep the notation and the discussion simple, assume that the identification number of site S_i is i and that the chosen coordinator will always be the active site with the largest identification number. Hence, when a coordinator fails, the algorithm must elect the active site that has the largest identification number. The algorithm must send this number to each active site in the system. In addition, the algorithm must provide a mechanism by which a site recovering from a crash can identify the current coordinator. Suppose that site S_i sends a request that is not answered by the coordinator within a prespecified time

interval T. In this situation, it is assumed that the coordinator has failed, and S_i tries to elect itself as the site for the new coordinator.

Site S_i sends an election message to every site that has a higher identification number. Site S_i then waits, for a time interval T, for an answer from any one of these sites. If it receives no response within time T, it assumes that all sites with numbers greater than i have failed, and it elects itself as the site for the new coordinator and sends a message to inform all active sites with identification numbers lower than i that it is the site at which the new coordinator resides.

If S_i does receive an answer, it begins a time interval T', to receive a message informing it that a site with a higher identification number has been elected. (Some other site is electing itself coordinator, and should report the results within time T'.) If S_i receives no message within T', then it assumes the site with a higher number has failed, and site S_i restarts the algorithm.

After a failed site recovers, it immediately begins execution of the same algorithm. If there are no active sites with higher numbers, the recovered site forces all sites with lower numbers to let it become the coordinator site, even if there is a currently active coordinator with a lower number. It is for this reason that the algorithm is termed the *bully* algorithm.

19.7 Distributed Query Processing

In Chapter 14, we saw that there are a variety of methods for computing the answer to a query. We examined several techniques for choosing a strategy for processing a query that minimize the amount of time that it takes to compute the answer. For centralized systems, the primary criterion for measuring the cost of a particular strategy is the number of disk accesses. In a distributed system, we must take into account several other matters, including

- The cost of data transmission over the network
- The potential gain in performance from having several sites process parts of the query in parallel

The relative cost of data transfer over the network and data transfer to and from disk varies widely depending on the type of network and on the speed of the disks. Thus, in general, we cannot focus solely on disk costs or on network costs. Rather, we must find a good tradeoff between the two.

19.7.1 Query Transformation

Consider an extremely simple query: "Find all the tuples in the *account* relation." Although the query is simple — indeed, trivial—processing it is not trivial, since the *account* relation may be fragmented, replicated, or both, as we saw in Section 19.2. If the *account* relation is replicated, we have a choice of replica to make. If no replicas are fragmented, we choose the replica for which the transmission cost is lowest. However, if a replica is fragmented, the choice is not so easy to make, since we need to compute several joins or unions to reconstruct the *account* relation. In this case,

the number of strategies for our simple example may be large. Query optimization by exhaustive enumeration of all alternative strategies may not be practical in such situations.

Fragmentation transparency implies that a user may write a query such as

$$\sigma_{branch\text{-}name = \text{“Hillside”}} (account)$$

Since *account* is defined as

$$account_1 \cup account_2$$

the expression that results from the name translation scheme is

$$\sigma_{branch\text{-}name = \text{“Hillside”}} (account_1 \cup account_2)$$

Using the query-optimization techniques of Chapter 13, we can simplify the preceding expression automatically. The result is the expression

$$\sigma_{branch\text{-}name = \text{“Hillside”}} (account_1) \cup \sigma_{branch\text{-}name = \text{“Hillside”}} (account_2)$$

which includes two subexpressions. The first involves only $account_1$, and thus can be evaluated at the Hillside site. The second involves only $account_2$, and thus can be evaluated at the Valleyview site.

There is a further optimization that can be made in evaluating

$$\sigma_{branch\text{-}name = \text{“Hillside”}} (account_1)$$

Since $account_1$ has only tuples pertaining to the Hillside branch, we can eliminate the selection operation. In evaluating

$$\sigma_{branch\text{-}name = \text{“Hillside”}} (account_2)$$

we can apply the definition of the $account_2$ fragment to obtain

$$\sigma_{branch\text{-}name = \text{“Hillside”}} (\sigma_{branch\text{-}name = \text{“Valleyview”}} (account))$$

This expression is the empty set, regardless of the contents of the *account* relation.

Thus, our final strategy is for the Hillside site to return $account_1$ as the result of the query.

19.7.2 Simple Join Processing

As we saw in Chapter 13, a major decision in the selection of a query-processing strategy is choosing a join strategy. Consider the following relational-algebra expression:

$$account \bowtie depositor \bowtie branch$$

Assume that the three relations are neither replicated nor fragmented, and that *account* is stored at site S_1, *depositor* at S_2, and *branch* at S_3. Let S_I denote the site at which the query was issued. The system needs to produce the result at site S_I. Among the possible strategies for processing this query are these:

- Ship copies of all three relations to site S_I. Using the techniques of Chapter 13, choose a strategy for processing the entire query locally at site S_I.
- Ship a copy of the *account* relation to site S_2, and compute $temp_1$ = *account* $\bowtie$ *depositor* at S_2. Ship $temp_1$ from S_2 to S_3, and compute $temp_2$ = $temp_1$ $\bowtie$ *branch* at S_3. Ship the result $temp_2$ to S_I.
- Devise strategies similar to the previous one, with the roles of S_1, S_2, S_3 exchanged.

No one strategy is always the best one. Among the factors that must be considered are the volume of data being shipped, the cost of transmitting a block of data between a pair of sites, and the relative speed of processing at each site. Consider the first two strategies listed. If we ship all three relations to S_I, and indices exist on these relations, we may need to re-create these indices at S_I. This re-creation of indices entails extra processing overhead and extra disk accesses. However, the second strategy has the disadvantage that a potentially large relation (*customer* $\bowtie$ *account*) must be shipped from S_2 to S_3. This relation repeats the address data for a customer once for each account that the customer has. Thus, the second strategy may result in extra network transmission compared to the first strategy.

19.7.3 Semijoin Strategy

Suppose that we wish to evaluate the expression $r_1 \bowtie r_2$, where r_1 and r_2 are stored at sites S_1 and S_2, respectively. Let the schemas of r_1 and r_2 be R_1 and R_2. Suppose that we wish to obtain the result at S_1. If there are many tuples of r_2 that do not join with any tuple of r_1, then shipping r_2 to S_1 entails shipping tuples that fail to contribute to the result. We want to remove such tuples before shipping data to S_1, particularly if network costs are high.

A possible strategy to accomplish all this is:

1. Compute $temp_1 \leftarrow \Pi_{R_1 \cap R_2}(r_1)$ at S_1.
2. Ship $temp_1$ from S_1 to S_2.
3. Compute $temp_2 \leftarrow r_2 \bowtie temp_1$ at S_2.
4. Ship $temp_2$ from S_2 to S_1.
5. Compute $r_1 \bowtie temp_2$ at S_1. The resulting relation is the same as $r_1 \bowtie r_2$.

Before considering the efficiency of this strategy, let us verify that the strategy computes the correct answer. In step 3, $temp_2$ has the result of $r_2 \bowtie \Pi_{R_1 \cap R_2}(r_1)$. In step 5, we compute

$$r_1 \bowtie r_2 \bowtie \Pi_{R_1 \cap R_2}(r_1)$$

Since join is associative and commutative, we can rewrite this expression as

$$(r_1 \bowtie \Pi_{R_1 \cap R_2}(r_1)) \bowtie r_2$$

Since $r_1 \bowtie \Pi_{(R_1 \cap R_2)}(r_1) = r_1$, the expression is, indeed, equal to $r_1 \bowtie r_2$, the expression we are trying to evaluate.

This strategy is particularly advantageous when relatively few tuples of r_2 contribute to the join. This situation is likely to occur if r_1 is the result of a relational-algebra expression involving selection. In such a case, $temp_2$ may have significantly fewer tuples than r_2. The cost savings of the strategy result from having to ship only $temp_2$, rather than all of r_2, to S_1. Additional cost is incurred in shipping $temp_1$ to S_2. If a sufficiently small fraction of tuples in r_2 contribute to the join, the overhead of shipping $temp_1$ will be dominated by the savings of shipping only a fraction of the tuples in r_2.

This strategy is called a **semijoin strategy**, after the semijoin operator of the relational algebra, denoted $\ltimes$. The semijoin of r_1 with r_2, denoted $r_1 \ltimes r_2$, is

$$\Pi_{R_1}(r_1 \bowtie r_2)$$

Thus, $r_1 \ltimes r_2$ selects those tuples of r_1 that contributed to $r_1 \bowtie r_2$. In step 3, $temp_2 = r_2 \ltimes r_1$.

For joins of several relations, this strategy can be extended to a series of semijoin steps. A substantial body of theory has been developed regarding the use of semijoins for query optimization. Some of this theory is referenced in the bibliographical notes.

19.7.4 Join Strategies that Exploit Parallelism

Consider a join of four relations:

$$r_1 \bowtie r_2 \bowtie r_3 \bowtie r_4$$

where relation r_i is stored at site S_i. Assume that the result must be presented at site S_1. There are many possible strategies for parallel evaluation. (We study the issue of parallel processing of queries in detail in Chapter 20.) In one such strategy, r_1 is shipped to S_2, and $r_1 \bowtie r_2$ computed at S_2. At the same time, r_3 is shipped to S_4, and $r_3 \bowtie r_4$ computed at S_4. Site S_2 can ship tuples of $(r_1 \bowtie r_2)$ to S_1 as they are produced, rather than wait for the entire join to be computed. Similarly, S_4 can ship tuples of $(r_3 \bowtie r_4)$ to S_1. Once tuples of $(r_1 \bowtie r_2)$ and $(r_3 \bowtie r_4)$ arrive at S_1, the computation of $(r_1 \bowtie r_2) \bowtie (r_3 \bowtie r_4)$ can begin, with the pipelined join technique of Section 13.7.2.2. Thus, computation of the final join result at S_1 can be done in parallel with the computation of $(r_1 \bowtie r_2)$ at S_2, and with the computation of $(r_3 \bowtie r_4)$ at S_4.

19.8 Heterogeneous Distributed Databases

Many new database applications require data from a variety of preexisting databases located in a heterogeneous collection of hardware and software environments. Manipulation of information located in a heterogeneous distributed database requires an additional software layer on top of existing database systems. This software layer is called a **multidatabase system**. The local database systems may employ different logical models and data-definition and data-manipulation languages, and may differ in their concurrency-control and transaction-management mechanisms. A multidatabase system creates the illusion of logical database integration without requiring physical database integration.

Full integration of heterogeneous systems into a homogeneous distributed database is often difficult or impossible:

- **Technical difficulties.** The investment in application programs based on existing database systems may be huge, and the cost of converting these applications may be prohibitive.
- **Organizational difficulties.** Even if integration is *technically* possible, it may not be *politically* possible, because the existing database systems belong to different corporations or organizations. In such cases, it is important for a multidatabase system to allow the local database systems to retain a high degree of **autonomy** over the local database and transactions running against that data.

For these reasons, multidatabase systems offer significant advantages that outweigh their overhead. In this section, we provide an overview of the challenges faced in constructing a multidatabase environment from the standpoint of data definition and query processing. Section 24.6 provides an overview of transaction management issues in multidatabases.

19.8.1 Unified View of Data

Each local database management system may use a different data model. For instance, some may employ the relational model, whereas others may employ older data models, such as the network model (see Appendix A) or the hierarchical model (see Appendix B).

Since the multidatabase system is supposed to provide the illusion of a single, integrated database system, a common data model must be used. A commonly used choice is the relational model, with SQL as the common query language. Indeed, there are several systems available today that allow SQL queries to a nonrelational database management system.

Another difficulty is the provision of a common conceptual schema. Each local system provides its own conceptual schema. The multidatabase system must integrate these separate schemas into one common schema. Schema integration is a complicated task, mainly because of the semantic heterogeneity.

Schema integration is not simply straightforward translation between data-definition languages. The same attribute names may appear in different local databases but with different meanings. The data types used in one system may not be supported by other systems, and translation between types may not be simple. Even for identical data types, problems may arise from the physical representation of data: One system may use ASCII, another EBCDIC; floating-point representations may differ; integers may be represented in *big-endian* or *little-endian* form. At the semantic level, an integer value for length may be inches in one system and millimeters in another, thus creating an awkward situation in which equality of integers is only an approximate notion (as is always the case for floating-point numbers). The same name may appear in different languages in different systems. For example, a system based in the United States may refer to the city "Cologne," whereas one in Germany refers to it as "Köln."

All these seemingly minor distinctions must be properly recorded in the common global conceptual schema. Translation functions must be provided. Indices must be annotated for system-dependent behavior (for example, the sort order of nonalphanumeric characters is not the same in ASCII as in EBCDIC). As we noted earlier, the alternative of converting each database to a common format may not be feasible without obsoleting existing application programs.

19.8.2 Query Processing

Query processing in a heterogeneous database can be complicated. Some of the issues are:

- Given a query on a global schema, the query may have to be translated into queries on local schemas at each of the sites where the query has to be executed. The query results have to be translated back into the global schema.

  The task is simplified by writing **wrappers** for each data source, which provide a view of the local data in the global schema. Wrappers also translate queries on the global schema into queries on the local schema, and translate results back into the global schema. Wrappers may be provided by individual sites, or may be written separately as part of the multidatabase system.

  Wrappers can even be used to provide a relational view of nonrelational data sources, such as Web pages (possibly with forms interfaces), flat files, hierarchical and network databases, and directory systems.

- Some data sources may provide only limited query capabilities; for instance, they may support selections, but not joins. They may even restrict the form of selections, allowing selections only on certain fields; Web data sources with form interfaces are an example of such data sources. Queries may therefore have to be broken up, to be partly performed at the data source and partly at the site issuing the query.

- In general, more than one site may need to be accessed to answer a given query. Answers retrieved from the sites may have to be processed to remove duplicates. Suppose one site contains *account* tuples satisfying the selection $balance < 100$, while another contains *account* tuples satisfying $balance > 50$. A query on the entire account relation would require access to both sites and removal of duplicate answers resulting from tuples with balance between 50 and 100, which are replicated at both sites.

- Global query optimization in a heterogeneous database is difficult, since the query execution system may not know what the costs are of alternative query plans at different sites. The usual solution is to rely on only local-level optimization, and just use heuristics at the global level.

Mediator systems are systems that integrate multiple heterogeneous data sources, providing an integrated global view of the data and providing query facilities on the global view. Unlike full-fledged multidatabase systems, mediator systems do not bother about transaction processing. (The terms mediator and multidatabase are of-

ten used in an interchangeable fashion, and systems that are called mediators may support limited forms of transactions.) The term **virtual database** is used to refer to multidatabase/mediator systems, since they provide the appearance of a single database with a global schema, although data exist on multiple sites in local schemas.

19.9 Directory Systems

Consider an organization that wishes to make data about its employees available to a variety of people in the organization; example of the kinds of data would include name, designation, employee-id, address, email address, phone number, fax number, and so on. In the precomputerization days, organizations would create physical directories of employees and distribute them across the organization. Even today, telephone companies create physical directories of customers.

In general, a directory is a listing of information about some class of objects such as persons. Directories can be used to find information about a specific object, or in the reverse direction to find objects that meet a certain requirement. In the world of physical telephone directories, directories that satisfy lookups in the forward direction are called **white pages**, while directories that satisfy lookups in the reverse direction are called **yellow pages**.

In today's networked world, the need for directories is still present and, if anything, even more important. However, directories today need to be available over a computer network, rather than in a physical (paper) form.

19.9.1 Directory Access Protocols

Directory information can be made available through Web interfaces, as many organizations, and phone companies in particular do. Such interfaces are good for humans. However, programs too, need to access directory information. Directories can be used for storing other types of information, much like file system directories. For instance, Web browsers can store personal bookmarks and other browser settings in a directory system. A user can thus access the same settings from multiple locations, such as at home and at work, without having to share a file system.

Several **directory access protocols** have been developed to provide a standardized way of accessing data in a directory. The most widely used among them today is the **Lightweight Directory Access Protocol** (**LDAP**).

Obviously all the types of data in our examples can be stored without much trouble in a database system, and accessed through protocols such as JDBC or ODBC. The question then is, why come up with a specialized protocol for accessing directory information? There are at least two answers to the question.

- First, directory access protocols are simplified protocols that cater to a limited type of access to data. They evolved in parallel with the database access protocols.
- Second, and more important, directory systems provide a simple mechanism to name objects in a hierarchical fashion, similar to file system directory names,

which can be used in a distributed directory system to specify what information is stored in each of the directory servers. For example, a particular directory server may store information for Bell Laboratories employees in Murray Hill, while another may store information for Bell Laboratories employees in Bangalore, giving both sites autonomy in controlling their local data. The directory access protocol can be used to obtain data from both directories, across a network. More importantly, the directory system can be set up to automatically forward queries made at one site to the other site, without user intervention.

For these reasons, several organizations have directory systems to make organizational information available online. As may be expected, several directory implementations find it beneficial to use relational databases to store data, instead of creating special-purpose storage systems.

19.9.2 LDAP: Lightweight Directory Access Protocol

In general a directory system is implemented as one or more servers, which service multiple clients. Clients use the application programmer interface defined by directory system to communicate with the directory servers. Directory access protocols also define a data model and access control.

The **X.500 directory access protocol**, defined by the International Organization for Standardization (ISO), is a standard for accessing directory information. However, the protocol is rather complex, and is not widely used. The **Lightweight Directory Access Protocol** (**LDAP**) provides many of the X.500 features, but with less complexity, and is widely used. In the rest of this section, we shall outline the data model and access protocol details of LDAP.

19.9.2.1 LDAP Data Model

In LDAP directories store **entries**, which are similar to objects. Each entry must have a **distinguished name (DN)**, which uniquely identifies the entry. A DN is in turn made up of a sequence of **relative distinguished names (RDNs)**. For example, an entry may have the following distinguished name.

cn=Silberschatz, ou=Bell Labs, o=Lucent, c=USA

As you can see, the distinguished name in this example is a combination of a name and (organizational) address, starting with a person's name, then giving the organizational unit (ou), the organization (o), and country (c). The order of the components of a distinguished name reflects the normal postal address order, rather than the reverse order used in specifying path names for files. The set of RDNs for a DN is defined by the schema of the directory system.

Entries can also have attributes. LDAP provides binary, string, and time types, and additionally the types tel for telephone numbers, and PostalAddress for addresses (lines separated by a "$" character). Unlike those in the relational model, attributes

are multivalued by default, so it is possible to store multiple telephone numbers or addresses for an entry.

LDAP allows the definition of **object classes** with attribute names and types. Inheritance can be used in defining object classes. Moreover, entries can be specified to be of one or more object classes. It is not necessary that there be a single most-specific object class to which an entry belongs.

Entries are organized into a **directory information tree (DIT)**, according to their distinguished names. Entries at the leaf level of the tree usually represent specific objects. Entries that are internal nodes represent objects such as organizational units, organizations, or countries. The children of a node have a DN containing all the RDNs of the parent, and one or more additional RDNs. For instance, an internal node may have a DN `c=USA`, and all entries below it have the value `USA` for the RDN `c`.

The entire distinguished name need not be stored in an entry; The system can generate the distinguished name of an entry by traversing up the DIT from the entry, collecting the `RDN=value` components to create the full distinguished name.

Entries may have more than one distinguished name—for example, an entry for a person in more than one organization. To deal with such cases, the leaf level of a DIT can be an **alias**, which points to an entry in another branch of the tree.

19.9.2.2 Data Manipulation

Unlike SQL, LDAP does not define either a data-definition language or a data manipulation language. However, LDAP defines a network protocol for carrying out data definition and manipulation. Users of LDAP can either use an application programming interface, or use tools provided by various vendors to perform data definition and manipulation. LDAP also defines a file format called **LDAP Data Interchange Format (LDIF)** that can be used for storing and exchanging information.

The querying mechanism in LDAP is very simple, consisting of just selections and projections, without any join. A query must specify the following:

- A base—that is, a node within a DIT—by giving its distinguished name (the path from the root to the node).
- A search condition, which can be a Boolean combination of conditions on individual attributes. Equality, matching by wild-card characters, and approximate equality (the exact definition of approximate equality is system dependent) are supported.
- A scope, which can be just the base, the base and its children, or the entire subtree beneath the base.
- Attributes to return.
- Limits on number of results and resource consumption.

The query can also specify whether to automatically dereference aliases; if alias dereferences are turned off, alias entries can be returned as answers.

One way of querying an LDAP data source is by using LDAP URLs. Examples of LDAP URLs are:

```
ldap:://aura.research.bell-labs.com/o=Lucent,c=USA
ldap:://aura.research.bell-labs.com/o=Lucent,c=USA??sub?cn=Korth
```

The first URL returns all attributes of all entries at the server with organization being Lucent, and country being USA. The second URL executes a search query (selection) cn=Korth on the subtree of the node with distinguished name o=Lucent, c=USA. The question marks in the URL separate different fields. The first field is the distinguished name, here o=Lucent,c=USA. The second field, the list of attributes to return, is left empty, meaning return all attributes. The third attribute, sub, indicates that the entire subtree is to be searched. The last parameter is the search condition.

A second way of querying an LDAP directory is by using an application programming interface. Figure 19.6 shows a piece of C code used to connect to an LDAP server and run a query against the server. The code first opens a connection to an LDAP server by ldap_open and ldap_bind. It then executes a query by ldap_search_s. The arguments to ldap_search_s are the LDAP connection handle, the DN of the base from which the search should be done, the scope of the search, the search condition, the list of attributes to be returned, and an attribute called attrsonly, which, if set to 1, would result in only the schema of the result being returned, without any actual tuples. The last argument is an output argument that returns the result of the search as an LDAPMessage structure.

The first **for** loop iterates over and prints each entry in the result. Note that an entry may have multiple attributes, and the second **for** loop prints each attribute. Since attributes in LDAP may be multivalued, the third **for** loop prints each value of an attribute. The calls ldap_msgfree and ldap_value_free free memory that is allocated by the LDAP libraries. Figure 19.6 does not show code for handling error conditions.

The LDAP API also contains functions to create, update, and delete entries, as well as other operations on the DIT. Each function call behaves like a separate transaction; LDAP does not support atomicity of multiple updates.

19.9.2.3 Distributed Directory Trees

Information about an organization may be split into multiple DITs, each of which stores information about some entries. The **suffix** of a DIT is a sequence of RDN=value pairs that identify what information the DIT stores; the pairs are concatenated to the rest of the distinguished name generated by traversing from the entry to the root. For instance, the suffix of a DIT may be o=Lucent, c=USA, while another may have the suffix o=Lucent, c=India. The DITs may be organizationally and geographically separated.

A node in a DIT may contain a **referral** to another node in another DIT; for instance, the organizational unit Bell Labs under o=Lucent, c=USA may have its own DIT, in which case the DIT for o=Lucent, c=USA would have a node ou=Bell Labs representing a referral to the DIT for Bell Labs.

Referrals are the key component that help organize a distributed collection of directories into an integrated system. When a server gets a query on a DIT, it may

```
#include <stdio.h>
#include <ldap.h>
main() {
        LDAP *ld;
        LDAPMessage *res, *entry;
        char *dn, *attr, *attrList[] = {"telephoneNumber", NULL};
        BerElement *ptr;
        int vals, i;
        ld = ldap_open("aura.research.bell-labs.com", LDAP_PORT);
        ldap_simple_bind(ld, "avi", "avi-passwd") ;
        ldap_search_s(ld, "o=Lucent, c=USA", LDAP_SCOPE_SUBTREE, "cn=Korth",
                        attrList, /*attrsonly*/ 0, &res);
        printf("found %d entries", ldap_count_entries(ld, res));
        for (entry=ldap_first_entry(ld, res); entry != NULL;
                                entry = ldap_next_entry(ld, entry)
        {
                dn = ldap_get_dn(ld, entry);
                printf("dn: %s", dn);
                ldap_memfree(dn);
                for (attr = ldap_first_attribute(ld, entry, &ptr);
                        attr ! NULL;
                        attr = ldap_next_attribute(ld, entry, ptr))
                {
                        printf("%s: ", attr);
                        vals = ldap_get_values(ld, entry, attr);
                        for (i=0; vals[i] != NULL; i++)
                                printf("%s, ", vals[i]);
                        ldap_value_free(vals);
                }
        }
        ldap_msgfree(res);
        ldap_unbind(ld);
}
```

Figure 19.6 Example of LDAP code in C.

return a referral to the client, which then issues a query on the referenced DIT. Access to the referenced DIT is transparent, proceeding without the user's knowledge. Alternatively, the server itself may issue the query to the referred DIT and return the results along with locally computed results.

The hierarchical naming mechanism used by LDAP helps break up control of information across parts of an organization. The referral facility then helps integrate all the directories in an organization into a single virtual directory.

Although it is not an LDAP requirement, organizations often choose to break up information either by geography (for instance, an organization may maintain a directory for each site where the organization has a large presence) or by organizational

structure (for instance, each organizational unit, such as department, maintains its own directory).

Many LDAP implementations support master–slave and multimaster replication of DITs, although replication is not part of the current LDAP version 3 standard. Work on standardizing replication in LDAP is in progress.

19.10 Summary

- A distributed database system consists of a collection of sites, each of which maintains a local database system. Each site is able to process local transactions: those transactions that access data in only that single site. In addition, a site may participate in the execution of global transactions; those transactions that access data in several sites. The execution of global transactions requires communication among the sites.
- Distributed databases may be homogeneous, where all sites have a common schema and database system code, or heterogeneous, where the schemas and system codes may differ.
- There are several issues involved in storing a relation in the distributed database, including replication and fragmentation. It is essential that the system minimize the degree to which a user needs to be aware of how a relation is stored.
- A distributed system may suffer from the same types of failure that can afflict a centralized system. There are, however, additional failures with which we need to deal in a distributed environment, including the failure of a site, the failure of a link, loss of a message, and network partition. Each of these problems needs to be considered in the design of a distributed recovery scheme.
- To ensure atomicity, all the sites in which a transaction T executed must agree on the final outcome of the execution. T either commits at all sites or aborts at all sites. To ensure this property, the transaction coordinator of T must execute a commit protocol. The most widely used commit protocol is the two-phase commit protocol.
- The two-phase commit protocol may lead to blocking, the situation in which the fate of a transaction cannot be determined until a failed site (the coordinator) recovers. We can use the three-phase commit protocol to reduce the probability of blocking.
- Persistent messaging provides an alternative model for handling distributed transactions. The model breaks a single transaction into parts that are executed at different databases. Persistent messages (which are guaranteed to be delivered exactly once, regardless of failures), are sent to remote sites to request actions to be taken there. While persistent messaging avoids the blocking problem, application developers have to write code to handle various types of failures.

- The various concurrency-control schemes used in a centralized system can be modified for use in a distributed environment.
  - In the case of locking protocols, the only change that needs to be incorporated is in the way that the lock manager is implemented. There are a variety of different approaches here. One or more central coordinators may be used. If, instead, a distributed lock-manager approach is taken, replicated data must be treated specially.
  - Protocols for handling replicated data include the primary-copy, majority, biased, and quorum-consensus protocols. These have different tradeoffs in terms of cost and ability to work in the presence of failures.
  - In the case of timestamping and validation schemes, the only needed change is to develop a mechanism for generating unique global timestamps.
  - Many database systems support lazy replication, where updates are propagated to replicas outside the scope of the transaction that performed the update. Such facilities must be used with great care, since they may result in nonserializable executions.
- Deadlock detection in a distributed lock-manager environment requires cooperation between multiple sites, since there may be global deadlocks even when there are no local deadlocks.
- To provide high availability, a distributed database must detect failures, reconfigure itself so that computation may continue, and recover when a processor or a link is repaired. The task is greatly complicated by the fact that it is hard to distinguish between network partitions or site failures.

  The majority protocol can be extended by using version numbers to permit transaction processing to proceed even in the presence of failures. While the protocol has a significant overhead, it works regardless of the type of failure. Less-expensive protocols are available to deal with site failures, but they assume network partitioning does not occur.
- Some of the distributed algorithms require the use of a coordinator. To provide high availability, the system must maintain a backup copy that is ready to assume responsibility if the coordinator fails. Another approach is to choose the new coordinator after the coordinator has failed. The algorithms that determine which site should act as a coordinator are called election algorithms.
- Queries on a distributed database may need to access multiple sites. Several optimization techniques are available to choose which sites need to be accessed. Based on fragmentation and replication, the techniques can use semijoin techniques to reduce data transfer.
- Heterogeneous distributed databases allow sites to have their own schemas and database system code. A multidatabase system provides an environment in which new database applications can access data from a variety of preexisting databases located in various heterogeneous hardware and software environments. The local database systems may employ different logical mod-

els and data-definition and data-manipulation languages, and may differ in their concurrency-control and transaction-management mechanisms. A multidatabase system creates the illusion of logical database integration, without requiring physical database integration.

- Directory systems can be viewed as a specialized form of database, where information is organized in a hierarchical fashion similar to the way files are organized in a file system. Directories are accessed by standardized directory access protocols such as LDAP.

  Directories can be distributed across multiple sites to provide autonomy to individual sites. Directories can contain referrals to other directories, which help build an integrated view whereby a query is sent to a single directory, and it is transparently executed at all relevant directories.