3. Concurrency Control for Transactions *Part One*

CSEP 545 Transaction Processing
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Outline

1. A Simple System Model
2. Serializability Theory
3. Synchronization Requirements
   for Recoverability
4. Two-Phase Locking
5. Preserving Transaction Handshakes
6. Implementing Two-Phase Locking
7. Deadlocks

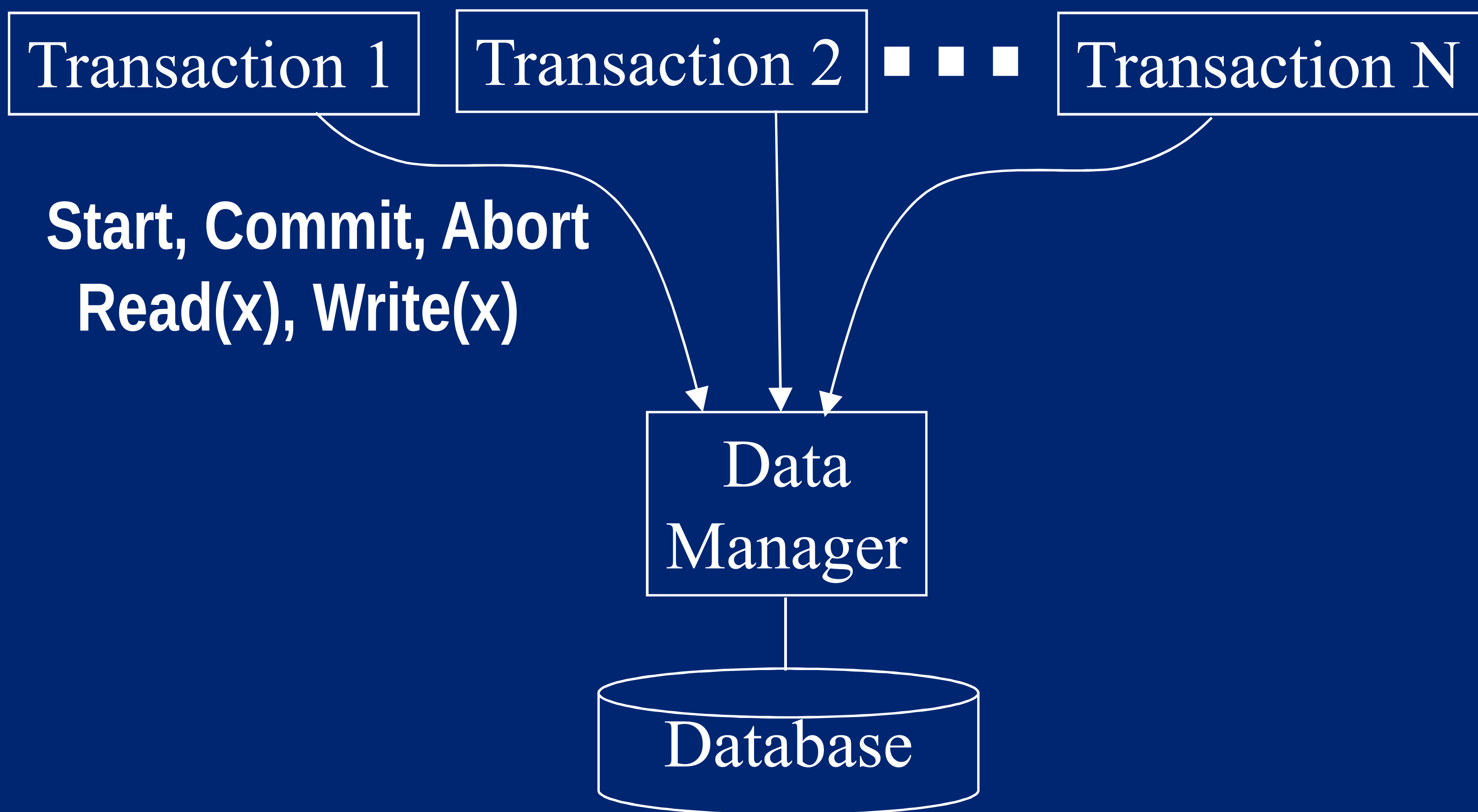
3.1 A Simple System Model

- Goal - Ensure serializable (SR) executions
- Implementation technique - Delay operations that would lead to non-SR results (e.g. set locks on shared data)
- For good performance minimize *overhead* and *delay* from synchronization operations
- First, we'll study how to get correct (SR) results
- Then, we'll study performance implications (mostly in Part Two)

Assumption - Atomic Operations

- We will synchronize Reads and Writes.
- We must therefore assume they're atomic
  - else we'd have to synchronize the finer-grained operations that implement Read and Write
- Read(x) - returns the current value of x in the DB
- Write(x, val) overwrites *all* of x (the *whole* page)
- This assumption of atomic operations is what allows us to abstract executions as sequences of reads and writes (without loss of information).
  - Otherwise, what would $w_k[x]$ $r_i[x]$ mean?
- Also, commit (c_i) and abort (a_i) are atomic

System Model



3.2 Serializability Theory

- The theory is based on modeling executions as histories, such as

$$H_1 = r_1[x]\ r_2[x]\ w_1[x]\ c_1\ w_2[y]\ c_2$$

- First, characterize a concurrency control algorithm by the properties of histories it allows.
- Then prove that any history having these properties is SR
- Why bother? It helps you understand why concurrency control algorithms work.

Equivalence of Histories

- Two operations conflict if their execution order affects their return values or the DB state.
  - a read and write on the same data item conflict
  - two writes on the same data item conflict
  - two reads (on the same data item) do *not* conflict
- Two histories are *equivalent* if they have the same operations and conflicting operations are in the same order in both histories
  - because only the relative order of conflicting operations can affect the result of the histories

Examples of Equivalence

- The following histories are equivalent

  $H_1 = r_1[x]\ r_2[x]\ w_1[x]\ c_1\ w_2[y]\ c_2$

  $H_2 = r_2[x]\ r_1[x]\ w_1[x]\ c_1\ w_2[y]\ c_2$

  $H_3 = r_2[x]\ r_1[x]\ w_2[y]\ c_2\ w_1[x]\ c_1$

  $H_4 = r_2[x]\ w_2[y]\ c_2\ r_1[x]\ w_1[x]\ c_1$
- But none of them are equivalent to

  $H_5 = r_1[x]\ w_1[x]\ r_2[x]\ c_1\ w_2[y]\ c_2$

  because $r_2[x]$ and $w_1[x]$ conflict and $r_2[x]$ precedes $w_1[x]$ in H_1 - H_4, but $r_2[x]$ follows $w_1[x]$ in H_5.

Serializable Histories

- A history is serializable if it is equivalent to a serial history
- For example,

  $H_1 = r_1[x]\ r_2[x]\ w_1[x]\ c_1\ w_2[y]\ c_2$

  is equivalent to

  $H_4 = r_2[x]\ w_2[y]\ c_2\ r_1[x]\ w_1[x]\ c_1$

  ($r_2[x]$ and $w_1[x]$ are in the same order in H_1 and H_4.)
- Therefore, H_1 is serializable.

Another Example

- $H_6 = r_1[x]\ r_2[x]\ w_1[x]\ r_3[x]\ w_2[y]\ w_3[x]\ c_3\ w_1[y]\ c_1\ c_2$ is equivalent to a serial execution of $T_2\ T_1\ T_3$,
  $H_7 = r_2[x]\ w_2[y]\ c_2\ r_1[x]\ w_1[x]\ w_1[y]\ c_1\ r_3[x]\ w_3[x]\ c_3$
- Each conflict implies a constraint on any equivalent serial history:

$T_2 \rightarrow T_3$

$H_6 = r_1[x]\ r_2[x]\ w_1[x]\ r_3[x]\ w_2[y]\ w_3[x]\ c_3\ w_1[y]\ c_1\ c_2$

$T_2 \rightarrow T_1$ $T_1 \rightarrow T_3$ $T_2 \rightarrow T_1$

Serialization Graphs

- A serialization graph, SG(H), for history H tells the effective execution order of transactions in H.
- Given history H, SG(H) is a directed graph whose nodes are the committed transactions and whose edges are all $T_i \rightarrow T_k$ such that at least one of T_i's operations precedes and conflicts with at least one of T_k's operations

$$H_6 = r_1[x]\ r_2[x]\ w_1[x]\ r_3[x]\ w_2[y]\ w_3[x]\ c_3\ w_1[y]\ c_1\ c_2$$

$$SG(H_6) = \quad \begin{tikzcd} T_2 \arrow[r] \arrow[rr, bend right] & T_1 \arrow[r] & T_3 \end{tikzcd}$$

The Serializability Theorem

A history is SR if and only if SG(H) is acyclic.

Proof: (if) SG(H) is acyclic. So let H_s be a serial history consistent with SG(H). Each pair of conflicting ops in H induces an edge in SG(H). Since conflicting ops in H_s and H are in the same order, $H_s \equiv H$, so H is SR.

(only if) H is SR. Let H_s be a serial history equivalent to H. We claim that if $T_i \rightarrow T_k$ in SG(H), then T_i precedes T_k in H_s (else $H_s \not\equiv H$). If SG(H) had a cycle, $T_1 \rightarrow T_2 \rightarrow \ldots \rightarrow T_n \rightarrow T_1$, then T_1 precedes T_1 in H_s, a contradiction. So SG(H) is acyclic.

How to Use the Serializability Theorem

- Characterize the set of histories that a concurrency control algorithm allows
- Prove that any such history must have an acyclic serialization graph.
- Therefore, the algorithm guarantees SR executions.
- We'll use this soon to prove that locking produces serializable executions.

3.3 Synchronization Requirements for Recoverability

- In addition to guaranteeing serializability, synchronization is needed to implement abort easily.
- When a transaction T aborts, the data manager wipes out all of T's effects, including
  - undoing T's writes that were applied to the DB, and
  - aborting transactions that read values written by T (these are called cascading aborts)
- Example - $w_1[x]\ r_2[x]\ w_2[y]$
  - to abort T_1, we must undo $w_1[x]$ *and* abort T_2 (a cascading abort)

Recoverability

- If T_k reads from T_i and T_i aborts, then T_k must abort
  - Example - $w_1[x]$ $r_2[x]$ a_1 implies T_2 must abort
- But what if T_k already committed? We'd be stuck.
  - Example - $w_1[x]$ $r_2[x]$ c_2 a_1
  - T_2 can't abort after it commits
- Executions must be *recoverable*:
  A transaction T's commit operation must follow the commit of every transaction from which T read.
  - Recoverable - $w_1[x]$ $r_2[x]$ c_1 c_2
  - Not recoverable - $w_1[x]$ $r_2[x]$ c_2 a_1
- Recoverability requires synchronizing operations.

Avoiding Cascading Aborts

- Cascading aborts are worth avoiding to
  - avoid complex bookkeeping, and
  - avoid an uncontrolled number of forced aborts
- To avoid cascading aborts, a data manager should ensure transactions only read committed data
- Example
  - avoids cascading aborts: $w_1[x]\ c_1\ r_2[x]$
  - allows cascading aborts: $w_1[x]\ r_2[x]\ a_1$
- A system that avoids cascading aborts also guarantees recoverability.

Strictness

- It's convenient to undo a write, w[x], by restoring its *before image* (=the value of x before w[x] executed)
- Example - $w_1[x,1]$ writes the value "1" into x.
  - $w_1[x,1]\ w_1[y,3]\ c_1\ w_2[y,1]\ r_2[x]\ a_2$
  - abort T_2 by restoring the before image of $w_2[y,1]$ (i.e. 3)
- But this isn't always possible.
  - For example, consider $w_1[x,2]\ w_2[x,3]\ a_1\ a_2$
  - a_1 & a_2 can't be implemented by restoring before images
  - notice that $w_1[x,2]\ w_2[x,3]\ a_2\ a_1$ would be OK
- A system is *strict* if it only reads or overwrites committed data.

Strictness (cont'd)

- More precisely, a system is *strict* if it only executes $r_i[x]$ or $w_i[x]$ if all previous transactions that wrote x committed or aborted.
- Examples ("…" marks a non-strict prefix)
  - strict: $w_1[x]\ c_1\ w_2[x]\ a_2$
  - not strict: $w_1[x]\ w_2[x]\ \ldots\ c_1\ a_2$
  - strict: $w_1[x]\ w_1[y]\ c_1\ r_2[x]\ w_2[y]\ a_2$
  - not strict: $w_1[x]\ w_1[y]\ r_2[x]\ \ldots\ c_1\ w_2[y]\ a_2$
  - To see why strictness matters in the above histories, consider what happens if T_1 aborts
- "Strict" implies "avoids cascading aborts."

3.4 Two-Phase Locking

- Basic locking - Each transaction sets a *lock* on each data item before accessing the data
  - the lock is a reservation
  - there are read locks and write locks
  - if one transaction has a write lock on x, then no other transaction can have any lock on x
- Example
  - $rl_i[x]$, $ru_i[x]$, $wl_i[x]$, $wu_i[x]$ denote lock/unlock operations
  - $wl_1[x]\ w_1[x]\ rl_2[x]\ r_2[x]$ is impossible
  - $wl_1[x]\ w_1[x]\ wu_1\mathbf{[x]}\ rl_2[x]\ r_2[x]$ is OK

Basic Locking Isn't Enough

- Basic locking doesn't guarantee serializability
- $rl_1[x]\ r_1[x]\ ru_1[x]$ $\qquad$ $wl_1[y]\ w_1[y]\ wu_1[y] c_1$

  $rl_2[y]\ r_2[y]\ wl_2[x]\ w_2[x]\ ru_2[y]\ wu_2[x]\ c_2$
- Eliminating the lock operations, we have
  $r_1[x]\ r_2[y]\ w_2[x]\ c_2\ w_1[y]\ c_1$ which isn't SR
- The problem is that locks aren't being released properly.

Two-Phase Locking (2PL) Protocol

- A transaction is *two-phase locked* if:
  - before reading x, it sets a read lock on x
  - before writing x, it sets a write lock on x
  - it holds each lock until after it executes the corresponding operation
  - after its first unlock operation, it requests no new locks
- Each transaction sets locks during a *growing phase* and releases them during a *shrinking phase*.
- Example - on the previous page T_2 is two-phase locked, but not T_1 since $ru_1[x] < wl_1[y]$
  - use "<" for "precedes"

2PL Theorem: If all transactions in an execution are two-phase locked, then the execution is SR.

Proof: Define $T_i \Rightarrow T_k$ if either

- T_i read x and T_k later wrote x, or
- T_i wrote x and T_k later read or wrote x

- If $T_i \Rightarrow T_k$, then T_i released a lock before T_k obtained some lock.
- If $T_i \Rightarrow T_k \Rightarrow T_m$, then T_i released a lock before T_m obtained some lock (because T_k is two-phase).
- If $T_i \Rightarrow \ldots \Rightarrow T_i$, then T_i released a lock before T_i obtained some lock, breaking the 2-phase rule.
- So there cannot be a cycle. By the Serializability Theorem, the execution is SR.

2PL and Recoverability

- 2PL does *not* guarantee recoverability
- This non-recoverable execution is 2-phase locked
  $wl_1[x]\ w_1[x]\ wu_1[x]\ rl_2[x]\ r_2[x]\ c_2\ \ldots\ c_1$
  - hence, it is not strict and allows cascading aborts
- However, holding write locks until *after* commit or abort guarantees strictness
  - and hence avoids cascading aborts and is recoverable
  - In the above example, T_1 must commit before its first unlock-write (wu_1): $wl_1[x]\ w_1[x]\ c_1\ wu_1[x]\ rl_2[x]\ r_2[x]\ c_2$

Automating Locking

- 2PL can be hidden from the application
- When a data manager gets a Read or Write operation from a transaction, it sets a read or write lock.
- How does the data manager know it's safe to release locks (and be two-phase)?
- Ordinarily, the data manager holds a transaction's locks until it commits or aborts. A data manager
  - can release <u>read</u> locks after it <u>receives</u> commit
  - releases <u>write</u> locks only after <u>processing</u> commit, to ensure strictness

3.5 Preserving Transaction Handshakes

- Read and Write are the only operations the system will control to attain serializability.
- So, if transactions communicate via messages, then implement SendMsg as Write, and ReceiveMsg as Read.
- Else, you could have the following:
  $w_1[x]\ r_2[x]\ send_2[M]\ receive_1[M]$
  - data manager didn't know about send/receive and thought the execution was SR.
- Also watch out for brain transport

Transactions Can Communicate via Brain Transport

```
T1:  Start
     . . .
     Display output
     Commit
```

User reads output

…

User enters input

Brain transport

```
T2:  Start
     Get input from display
     . . .
     Commit
```

Brain Transport (cont'd)

- For practical purposes, if user waits for T_1 to commit before starting T_2, then the data manager can ignore brain transport.
- This is called a transaction handshake (T_1 commits before T_2 starts)
- Reason - Locking preserves the order imposed by transaction handshakes
  - e.g., it serializes T_1 before T_2.

2PL Preserves Transaction Handshakes

- 2PL serializes transactions (abbr. txns) consistent with all transaction handshakes. I.e. there's an equivalent serial execution that preserves the transaction order of transaction handshakes
- This isn't true for arbitrary SR executions. E.g.
  - $r_1[x]\ w_2[x]\ c_2\ r_3[y]\ c_3\ w_1[y]\ c_1$
  - T_2 commits before T_3 starts, but the only equivalent serial execution is $T_3\ T_1\ T_2$
  - $rl_1[x]\ r_1[x]\ wl_1[y]\ ru_1[x]\ wl_2[x]\ w_2[x]\ wu_2[x]\ c_2$ but now we're stuck, since we can't set $rl_3[y]\ r_3[y]$. So the history cannot occur using 2PL.

2PL Preserves Transaction Handshakes (cont'd)

- Stating this more formally …
- Theorem:

  For any 2PL execution H,

  there is an equivalent serial execution H_s,

  such that for all T_i, T_k,

  if T_i committed before T_k started in H,

  then T_i precedes T_k in H_s.

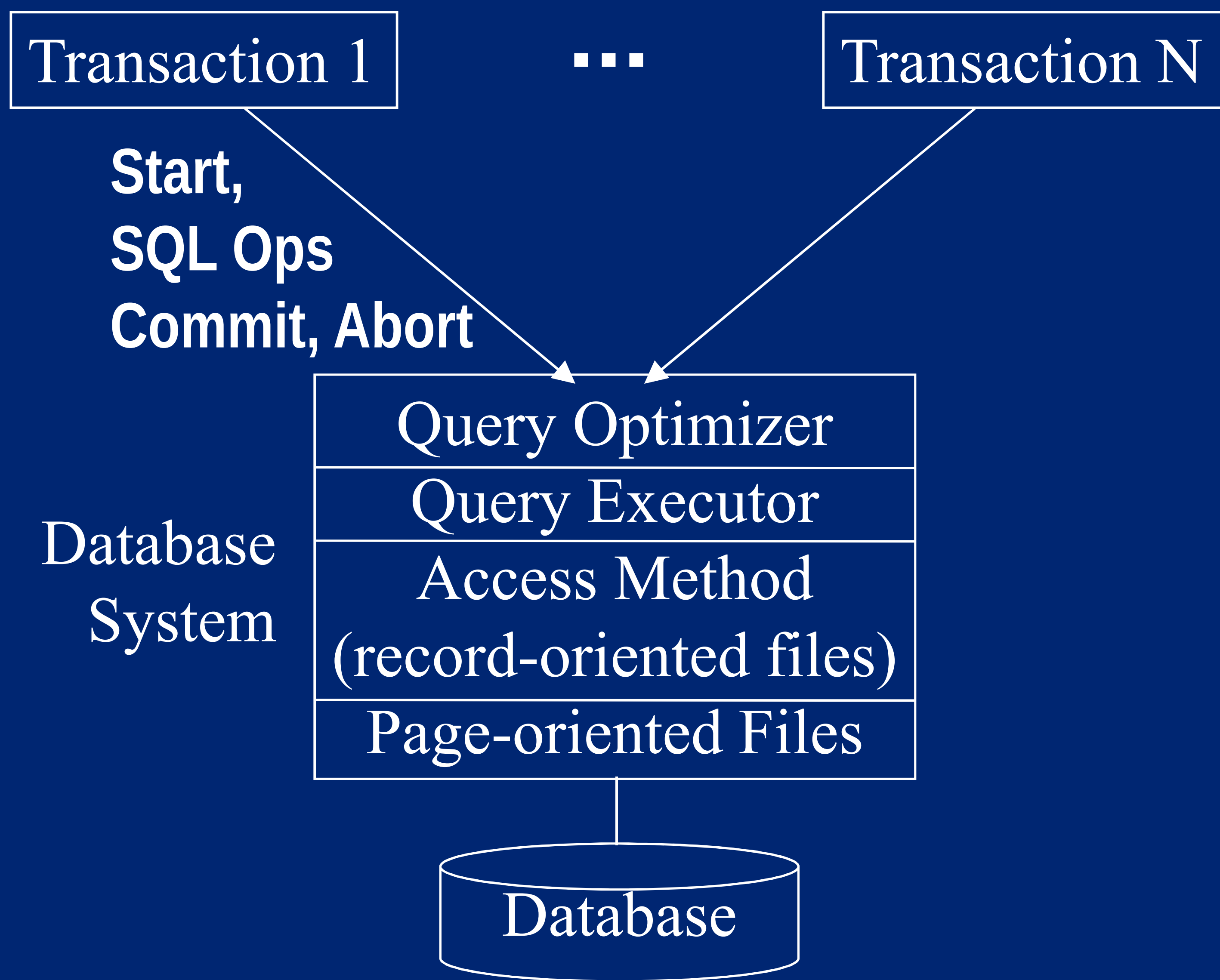
Brain Transport — One Last Time

- If a user reads committed displayed output of T_i and uses that displayed output as input to transaction T_k, then he/she should wait for T_i to commit before starting T_k.
- The user can then rely on transaction handshake preservation to ensure T_i is serialized before T_k.

3.6 Implementing Two-Phase Locking

- Even if you never implement a DB system, it’s valuable to understand locking implementation, because it can have a big effect on performance.
- A data manager implements locking by
  – implementing a lock manager
  – setting a lock for each Read and Write
  – handling deadlocks

System Model



How to Implement SQL

- Query Optimizer - translates SQL into an ordered expression of relational DB operators (Select, Project, Join)
- Query Executor - executes the ordered expression by running a program for each operator, which in turn accesses records of files
- Access methods - provides indexed record-at-a-time access to files (OpenScan, GetNext, …)
- Page-oriented files - Read or Write (page address)

Which Operations Get Synchronized?

Query Optimizer	← SQL operations
Query Executor	← SQL operations
Access Method (record-oriented files)	← Record-oriented operations
Page-oriented Files	← Page-oriented operations

- It's a tradeoff between
  - amount of concurrency and
  - runtime expense and programming complexity of synchronization

Lock Manager

- A lock manager services the operations
  - Lock(trans-id, data-item-id, mode)
  - Unlock(trans-id, data-item-id)
  - Unlock(trans-id)
- It stores locks in a lock table. Lock op inserts [trans-id, mode] in the table. Unlock deletes it.

Data Item	List of Locks	Wait List
x	[T_1,r] [T_2,r]	[T_3,w]
y	[T_4,w]	[T_5,w] [T_6, r]
⋮		

Lock Manager (cont'd)

- Caller generates data-item-id, e.g. by hashing data item name
- The lock table is hashed on data-item-id
- Lock and Unlock must be atomic, so access to the lock table must be "locked"
- Lock and Unlock are called frequently. They must be *very* fast. Average < 100 instructions.
  - This is hard, in part due to slow compare-and-swap operations needed for atomic access to lock table

Lock Manager (cont'd)

- In MS SQL Server
  - Locks are approx 32 bytes each.
  - Each lock contains a Database-ID, Object-Id, and other resource-specific lock information such as record id (RID) or key.
  - Each lock is attached to lock resource block (64 bytes) and lock owner block (32 bytes)

Locking Granularity

- <u>Granularity</u> - size of data items to lock
  - e.g., files, pages, records, fields
- Coarse granularity implies
  - very few locks, so little locking overhead
  - must lock large chunks of data, so high chance of conflict, so concurrency may be low
- Fine granularity implies
  - many locks, so high locking overhead
  - locking conflict occurs only when two transactions try to access the exact same data concurrently
- High performance TP requires record locking

Multigranularity Locking (MGL)

- Allow different txns to lock at different granularity
  - big queries should lock coarse-grained data (e.g. tables)
  - short transactions lock fine-grained data (e.g. rows)
- Lock manager can't detect these conflicts
  - each data item (e.g., table or row) has a different id
- Multigranularity locking "trick"
  - exploit the natural hierarchy of data containment
  - before locking fine-grained data, set *intention locks* on coarse grained data that contains it
  - e.g., before setting a read-lock on a row, get an intention-read-lock on the table that contains the row
  - Intention-read-locks conflicts with a write lock

3.7 Deadlocks

- A set of transactions is <u>deadlocked</u> if every transaction in the set is blocked and will remain blocked unless the system intervenes.
  - Example

	$rl_1[x]$	granted
	$rl_2[y]$	granted
	$wl_2[x]$	blocked
	$wl_1[y]$	blocked and deadlocked

- Deadlock is 2PL's way to avoid non-SR executions
  - $rl_1[x]\ r_1[x]\ rl_2[y]\ r_2[y]$ … can't run $w_2[x]\ w_1[y]$ and be SR
- To repair a deadlock, you <u>must</u> abort a transaction
  - if you released a transaction's lock without aborting it, you'd break 2PL

Deadlock Prevention

- Never grant a lock that can lead to deadlock
- Often advocated in operating systems
- Useless for TP, because it would require running transactions serially.
  - Example to prevent the previous deadlock, $rl_1[x]\ rl_2[y]\ wl_2[x]\ wl_1[y]$, the system can't grant $rl_2[y]$
- Avoiding deadlock by resource ordering is unusable in general, since it overly constrains applications.
  - But may help for certain high frequency deadlocks
- Setting all locks when txn begins requires too much advance knowledge and reduces concurrency.

Deadlock Detection

- Detection approach: Detect deadlocks automatically, and abort a deadlocked transactions (the <u>victim</u>).
- It's the preferred approach, because it
  - allows higher resource utilization and
  - uses cheaper algorithms
- Timeout-based deadlock detection - If a transaction is blocked for too long, then abort it.
  - Simple and easy to implement
  - But aborts unnecessarily and
  - some deadlocks persist for too long

Detection Using Waits-For Graph

- Explicit deadlock detection - Use a Waits-For Graph
  - Nodes = {transactions}
  - Edges = $\{T_i \rightarrow T_k \mid T_i$ is waiting for T_k to release a lock$\}$
  - Example (previous deadlock) $T_1 \leftrightarrows T_2$
- Theorem: If there's a deadlock, then the waits-for graph has a cycle.

Detection Using Waits-For Graph (cont'd)

- So, to find deadlocks
  – when a transaction blocks, add an edge to the graph
  – periodically check for cycles in the waits-for graph
- Need not test for deadlocks too often. (A cycle won't disappear until you detect it and break it.)
- When a deadlock is detected, select a victim from the cycle and abort it.
- Select a victim that hasn't done much work (e.g., has set the fewest locks).

Cyclic Restart

- Transactions can cause each other to abort forever.
  - T_1 starts running. Then T_2 starts running.
  - They deadlock and T_1 (the oldest) is aborted.
  - T_1 restarts, bumps into T_2 and again deadlocks
  - T_2 (the oldest) is aborted ...
- Choosing the youngest in a cycle as victim avoids cyclic restart, since the oldest running transaction is never the victim.
- Can combine with other heuristics, e.g. fewest-locks

MS SQL Server

- Aborts the transaction that is “cheapest” to roll back.
  - “Cheapest” is determined by the amount of log generated.
  - Allows transactions that you’ve invested a lot in to complete.
- SET DEADLOCK_PRIORITY LOW (vs. NORMAL) causes a transaction to sacrifice itself as a victim.

Distributed Locking

- Suppose a transaction can access data at many data managers
- Each data manager sets locks in the usual way
- When a transaction commits or aborts, it runs two-phase commit to notify all data managers it accessed
- The only remaining issue is distributed deadlock

Distributed Deadlock

- The deadlock spans two nodes.
  Neither node alone can see it.

Node 1	Node 2
$rl_1[x]$	$rl_2[y]$
$wl_2[x]$ (blocked)	$wl_1[y]$ (blocked)

- Timeout-based detection is popular. Its weaknesses are less important in the distributed case:
  - aborts unnecessarily and some deadlocks persist too long
  - possibly abort younger unblocked transaction to avoid cyclic restart

Oracle Deadlock Handling

- Uses a waits-for graph for single-server deadlock detection.
- The transaction that detects the deadlock is the victim.
- Uses timeouts to detect distributed deadlocks.

Fancier Dist'd Deadlock Detection

- Use waits-for graph cycle detection with a central deadlock detection server
  - more work than timeout-based detection, and no evidence it does better, performance-wise
  - phantom deadlocks? - No, because each waits-for edge is an SG edge. So, WFG cycle => SG cycle (modulo spontaneous aborts)
- Path pushing (a.k.a. flooding) - Send paths $T_i \rightarrow \ldots \rightarrow T_k$ to each node where T_k might be blocked.
  - Detects short cycles quickly
  - Hard to know where to send paths. Possibly too many messages

What's Coming in Part Two?

- Locking Performance
- More details on multigranularity locking
- Hot spot techniques
- Query-Update Techniques
- Phantoms
- B-Trees and Tree locking

Locking Performance

- The following is oversimplified. We'll revisit it.
- Deadlocks are rare.
  - Typically 1-2% of transactions deadlock.
- Locking performance problems are *not* rare.
- The problem is too much blocking.
- The solution is to reduce the "locking load"
- Good heuristic – If more than 30% of transactions are blocked, then reduce the number of concurrent transactions