5. Concurrency Control for Transactions

CSE 593 Transaction Processing
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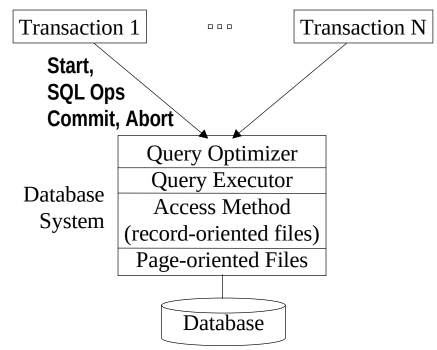
Outline

1. A Model for Concurrency Control
2. Serializability Theory
3. Synchronization Requirements for Recoverability
4. Two-Phase Locking
5. Implementing Two-Phase Locking
6. Locking Performance
7. Hot Spot Techniques
8. Query-Update Techniques
9. Phantoms
10. B-Trees
11. Tree locking

5.1 A Model for Concurrency Control The Problem

- 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

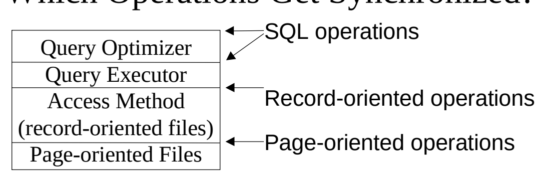
System Model



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?



- It's a tradeoff between
  - amount of concurrency and
  - overhead and complexity of synchronization
- For now, assume page operations
  - notation: $r_i[x]$, $w_i[x]$ where "x" is a page

and use the neutral term <u>data manager</u>

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?

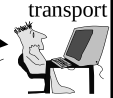
Assumption - Txns communicate *only* via Read and Write

- 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




```
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.
- Stating this precisely and proving it is non-trivial.
- … more later ….

5.2 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 $w_1[x]$ precedes $r_2[x]$ 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:

$T_2 \rightarrow T_3$

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

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

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

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

$SG(H_6) = \quad T_2 \rightarrow T_1 \rightarrow T_3$

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. 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.

5.3 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 $w_2[y,1]$, = 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: $w_1[x]\ w_2[x]\ \ldots\ a_1\ a_2$
  - strict: $w_1[x]\ w_1[y]\ c_1\ w_2[y]\ r_2[x]\ a_2$
  - not strict: $w_1[x]\ w_1[y]\ w_2[y]\ a_1\ r_2[x]\ a_2$
- "Strict" implies "avoids cascading aborts."

5.4 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
  - $wl_1[x]\ w_1[x]\ wu_1$**[x]** $rl_2[x]\ r_2[x]$ is OK

Basic Locking Isn't Enough

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

  $rl_2[y]\ r_2[y]\ wl_2[x]\ w_2[x]\ ru_2[y]\ wu_2[x]\ c_2$
- 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 $T_i \Rightarrow ... \Rightarrow T_i$, 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 it's 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

2PL Preserves Transaction Handshakes

- Recall the definition: T_i commits before T_k starts
- 2PL serializes 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 $T_3\ T_1\ T_2$
  - $rl_1[x]\ r_1[x]\ wl_1[y]\ ru_1[x]\ wl_2[x]\ w_2[x]\ wu_2[x]\ c_2$ (stuck, can't set $rl_3[y]$) $r_3[y]$ … so not 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.

5.5 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

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)

Deadlocks

- A set of transactions is deadlocked 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 must 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 victim).
- 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
- Don't 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 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 - 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

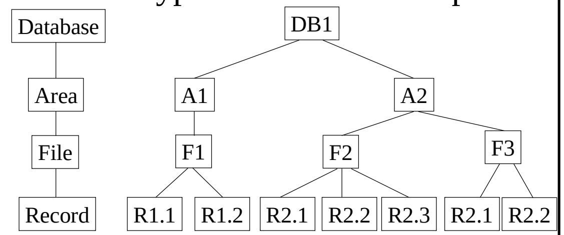
Locking Granularity

- Granularity - 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

MGL Type and Instance Graphs



Lock Type Graph

Lock Instance Graph

- Before setting a read lock on R2.3, first set an intention-read lock on DB1, then A2, and then F2.
- Set locks root-to-leaf. Release locks leaf-to-root.

MGL Compatibility Matrix

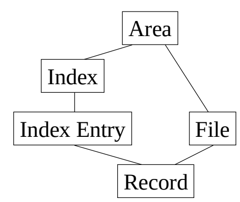
	r	w	ir	iw	riw
r	y	n	y	n	n
w	n	n	n	n	n
ir	y	n	y	y	y
iw	n	n	y	y	n
riw	n	n	y	n	n

riw = read with intent to write, for a scan that updates some of the records it reads

- E.g., ir conflicts with w because ir says there's a fine-grained r-lock that conflicts with a w-lock on the container
- To r-lock an item, need an r-, ir- or riw-lock on its parent
- To w-lock an item, need a w-, iw- or riw-lock on its parent

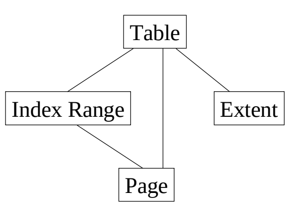
MGL Complexities

- Relational DBMSs use MGL to lock SQL queries, short updates, and scans with updates.
- Use lock escalation - start locking at fine-grain and escalate to coarse grain after n^{th} lock is set.
- The lock type graph is a directed acyclic graph, not a tree, to cope with indices
- R-lock one path to an item. W-lock all paths to it.



MS SQL Server

- MS SQL Server can lock at table, page, and row level.
- Uses intention read ("share") and intention write ("exclusive") locks at the table and page level.
- Tries to avoid escalation by choosing the "appropriate" granularity when the scan is instantiated.



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5.6 Locking Performance

- Deadlocks are rare
  - up to 1% - 2% of transactions deadlock
- The one exception to this is lock conversions
  - r-lock a record and later upgrade to w-lock
  - e.g., T_i = read(x) … write(x)
  - if two txns do this concurrently, they'll deadlock (both get an r-lock on x before either gets a w-lock)
  - To avoid lock conversion deadlocks, get a w-lock first and down-grade to an r-lock if you don't need to write.
  - Use SQL Update statement or explicit program hints

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Conversions in MS SQL Server

- Update-lock prevents lock conversion deadlock.
  - Conflicts with other update and write locks, but not with read locks.
  - Only on pages and rows (not tables)
- You get an update lock by using the UPDLOCK hint in the FROM clause

```
Select Foo.A
From Foo (UPDLOCK)
Where Foo.B = 7
```

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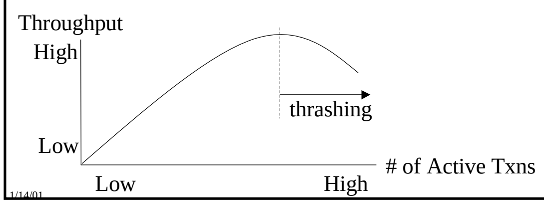
Blocking and Lock Thrashing

- The locking performance problem is too much delay due to blocking
  - little delay until locks are saturated
  - then major delay, due to the locking bottleneck
  - thrashing - the point where throughput decreases with increasing load



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More on Thrashing

- It's purely a blocking problem
  - It happens even when the abort rate is low
- As number of transactions increase
  - each additional transaction is more likely to block
  - but first, it gathers some locks, increasing the probability others will block (negative feedback)

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Avoiding Thrashing

- If over 30% of active transactions are blocked, then the system is (nearly) thrashing so reduce the number of active transactions
- Timeout-based deadlock detection mistakes
  - They happen due to long lock delays
  - So the system is probably close to thrashing
  - So if deadlock detection rate is too high (over 2%) reduce the number of active transactions

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Interesting Sidelights

- By getting all locks before transaction Start, you can increase throughput at the thrashing point because blocked transactions hold no locks
  - But it assumes you get exactly the locks you need and retries of get-all-locks are cheap
- Pure restart policy - abort when there's a conflict and restart when the conflict disappears
  - If aborts are cheap and there's low contention for other resources, then this policy produces higher throughput before thrashing than a blocking policy
  - But response time is greater than a blocking policy

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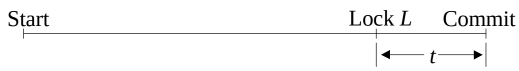
How to Reduce Lock Contention

- If each transaction holds a lock *L* for *t* seconds, then the maximum throughput is 1/*t* txns/second



- To increase throughput, reduce *t* (lock holding time)
  - Set the lock later in the transaction's execution (e.g., defer updates till commit time)
  - Reduce transaction execution time (reduce path length, read from disk before setting locks)
  - Split a transaction into smaller transactions

Reducing Lock Contention (cont'd)

- Reduce number of conflicts
  - Use finer grained locks, e.g., by partitioning tables vertically

Part#	Price	OnHand	PartName	CatalogPage

Part#	Price	OnHand

Part#	PartName	CatalogPage

Mathematical Model of Locking

- K locks per transaction
- N transactions
- D lockable data items
- T time between lock requests

- N transactions each own K/2 locks on average
  - KN/2 in total
- Each lock request has probability KN/2D of conflicting with an existing lock.
- Each transaction requests K locks, so its probability of experiencing a conflict is $K^2N/2D$.
- Probability of a deadlock is proportional to K^4N/D^2
  - Prob(deadlock) / Prop(conflict) = K^2/D
  - if K=10 and D = 10^6, then K^2/D = .0001

5.7 Hot Spot Techniques

- If each txn holds a lock for *t* seconds, then the max throughput is 1/*t* txns/second for that lock.
- Hot spot - A data item that's more popular than others, so a large fraction of active txns need it
  - Summary information (total inventory)
  - End-of-file marker in data entry application
  - Counter used for assigning serial numbers
- Hot spots often create a convoy of transactions. The hot spot lock serializes transactions.

Hot Spot Techniques (cont'd)

- Special techniques are needed to reduce *t*
  - Keep the hot data in main memory
  - Delay operations on hot data till commit time
  - Use optimistic methods
  - Batch up operations to hot spot data
  - Partition hot spot data

Delaying Operations Until Commit

- Data manager logs each transaction's updates
- Only applies the updates (and sets locks) after receiving Commit from the transaction
- IMS Fast Path uses this for
  - Data Entry DB
  - Main Storage DB
- Works for write, insert, and delete, but not read

Locking Higher-Level Operations

- Read is often part of a read-write pair, such as Increment(x, n), which adds constant n to x, but doesn't return a value.
- Increment (and Decrement) commute
- So, introduce Increment and Decrement locks

	r	w	inc	dec
r	y	n	n	n
w	n	n	n	n
inc	n	n	y	y
dec	n	n	y	y

- But if Inc and Dec have a threshold (e.g. a quantity of zero), then they conflict (when the threshold is near)

Solving the Threshold Problem

Another IMS Fast Path Technique

- Use a blind Decrement (no threshold) and Verify(x, n), which returns true if $x \geq n$
- Re-execute Verify at commit time
  - If it returns a different value than it did during normal execution, then abort
  - It's like checking that the threshold lock you didn't set during Decrement is still valid.

```
bEnough = Verify(iQuantity, n);
If (bEnough) Decrement(iQuantity, n)
else print ("not enough");
```


Optimistic Concurrency Control

- The Verify trick is optimistic concurrency control
- Main idea - execute operations on shared data without setting locks. At commit time, test if there were conflicts on the locks (that you didn't set).
- Often used in client/server systems
  - Client does all updates in cache without shared locks
  - At commit time, try to get locks and perform updates

Batching

- Transactions add updates to a mini-batch and only periodically apply the mini-batch to shared data.
  - Each process has a private data entry file, in addition to a global shared data entry file
  - Each transaction appends to its process' file
  - Periodically append the process file to the shared file
- Tricky failure handling
  - Gathering up private files
  - Avoiding holes in serial number order

Partitioning

- Split up inventory into partitions
- Each transaction only accesses one partition
- Example
  - Each ticket agency has a subset of the tickets
  - If one agency sells out early, it needs a way to get more tickets from other agencies (partitions)

5.8 Query-Update Techniques

- Queries run for a long time and lock a lot of data — a performance nightmare when trying also to run short update transactions
- There are several good solutions
  - Use a data warehouse
  - Accept weaker consistency guarantees
  - Use multiversion data
- Solutions trade data quality or timeliness for performance

Data Warehouse

- A data warehouse contains a snapshot of the DB which is periodically refreshed from the TP DB
- All queries run on the data warehouse
- All update transactions run on the TP DB
- Queries don't get absolutely up-to-date data
- How to refresh the data warehouse?
  - Stop processing transactions and copy the TP DB to the data warehouse. Possibly run queries while refreshing
  - Treat the warehouse as a DB replica and use a replication technique

Degrees of Isolation

- Serializability = *Degree 3 Isolation*
- Degree 2 Isolation (a.k.a. cursor stability)
  - Data manager holds read-lock(x) only while reading x, but holds write locks till commit (as in 2PL)
  - E.g. when scanning records in a file, each get-next-record releases lock on current record and gets lock on next one
  - read(x) is not "repeatable" within a transaction, e.g., $rl_1[x]\ r_1[x]\ ru_1[x]\ wl_2[x]\ w_2[x]\ wu_2[x]\ rl_1[x]\ r_1[x]\ ru_1[x]$
  - Degree 2 is commonly used by ISAM file systems
  - Degree 2 is often a DB system's default behavior! And customers seem to accept it!!!

Degrees of Isolation (cont'd)

- Could run queries Degree 2 and updaters Degree 3
  - Updaters are still serializable w.r.t. each other
- Degree 1 - no read locks; hold write locks to commit
- Unfortunately, SQL concurrency control standards have been stated in terms of "repeatable reads" and "cursor stability" instead of serializability, leading to much confusion.

ANSI SQL Isolation Levels

- Uncommitted Read - Degree 1
- Committed Read - Degree 2
- Repeatable Read - Uses read locks and write locks, but allows "phantoms"
- Serializable - Degree 3

MS SQL Server

- Lock hints in SQL FROM clause
  - All the ANSI isolation levels, plus …
  - UPDLOCK - use update locks instead of read locks
  - READPAST - ignore locked rows (if running read committed)
  - PAGLOCK - use page lock when the system would otherwise use a table lock
  - TABLOCK - shared table lock till end of command or transaction
  - TABLOCKX - exclusive table lock till end of command or transaction

Multiversion Data

- Assume record granularity locking
- Each write operation creates a new version instead of overwriting existing value.
- So each logical record has a sequence of versions.
- Tag each record with transaction id of the transaction that wrote that version

Tid	Previous	E#	Name	Other fields
123	null	1	Bill	
175	123	1	Bill	
134	null	2	Sue	
199	134	2	Sue	
227	null	27	Steve	

Multiversion Data (cont'd)

- Execute update transactions using ordinary 2PL
- Execute queries in *snapshot mode*
  - System keeps a <u>commit list</u> of tids of all committed txns
  - When a query starts executing, it reads the commit list
  - When a query reads x, it reads the latest version of x written by a transaction on its commit list
  - Thus, it reads the database state that existed when it started running

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Commit List Management

- Maintain and periodically recompute a tid T-Oldest, such that
  - Every active txn's tid is greater than T-Oldest
  - Every new tid is greater than T-Oldest
  - For every committed transaction with tid ≤ T-Oldest, its versions are committed
  - For every aborted transaction with tid ≤ T-Oldest, its versions are wiped out
- Queries don't need to know tids ≤ T-Oldest
  - So only maintain the commit list for tids > T-Oldest

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Multiversion Garbage Collection

- Can delete an old version of x if no query will ever read it
  - There's a later version of x whose tid ≥ T-Oldest (or is on every active query's commit list)
- Originally used in Prime Computer's CODASYL DB system and Oracle's Rdb/VMS

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Oracle Multiversion Concurrency Control

- Data page contains latest version of each record, which points to older version in rollback segment.
- Read-committed query reads data as of its start time.
- Read-only isolation reads data as of transaction start time.
- "Serializable" query reads data as of the txn's start time.
  - An update checks that the updated record was not modified after txn start time.
  - If that check fails, Oracle returns an error.
  - If there isn't enough history for Oracle to perform the check, Oracle returns an error. (You can control the history area's size.)
  - What if T_1 and T_2 modify each other's readset concurrently?

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Oracle Concurrency Control (cont'd)

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

- The result is not serializable!
- In any SR execution, one transaction would have read the other's output

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5.9 Phantoms

- Problems when using 2PL with inserts and deletes

Accounts

Acct#	Location	Balance
1	Seattle	400
2	Tacoma	200
3	Tacoma	300

Assets

Location	Total
Seattle	400
Tacoma	500

T_1: Read Accounts 1, 2, and 3
T_2: Insert Accounts[4, Tacoma, 100] ← The phantom record
T_2: Read Assets(Tacoma), returns 500
T_2: Write Assets(Tacoma, 600)
T_1: Read Assets(Tacoma), returns 600
T_1: Commit

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The Phantom Phantom Problem

- It looks like T_1 should lock record 4, which isn't there!
- Which of T_1's operations determined that there were only 3 records?
  - Read end-of-file?
  - Read record counter?
  - SQL Select operation?
- This operation conflicts with T_2's Insert Accounts[4,Tacoma,100]
- Therefore, Insert Accounts[4,Tacoma,100] shouldn't run until after T_1 commits

Avoiding Phantoms - Predicate Locks

- Suppose a query reads all records satisfying predicate P. For example,
  - Select * From Accounts Where Location = "Tacoma"
  - Normally would hash each record id to an integer lock id
  - And lock control structures. Too coarse grained.
- Ideally, set a read lock on P
  - which conflicts with a write lock Q if some record can satisfy (P and Q)
- For arbitrary predicates, this is too slow to check
  - Not within a few hundred instructions, anyway

Precision Locks

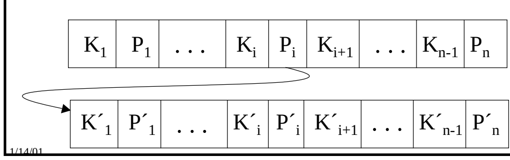
- Suppose update operations are on single records
- Maintain a list of predicate Read-locks
- Insert, Delete, & Update write-lock the record and check for conflict with all predicate locks
- Query sets a read lock on the predicate and check for conflict with all record locks
- Cheaper than predicate satisfiability, but still too expensive for practical implementation.

5.10 B-Trees

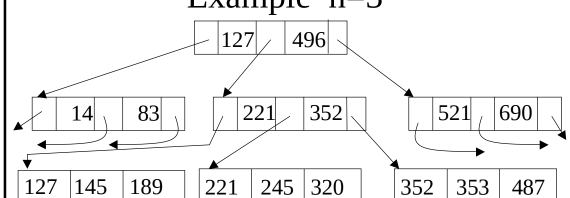
- An *index* maps field values to record ids.
  - Record id = [page-id, offset-within-page]
  - Most common DB index structures: hashing and B-trees
  - DB index structures are *page-oriented*
- Hashing uses a function $H:V \rightarrow B$, from field values to block numbers.
  - V = social security numbers. B = {1 .. 1000}
    H(v) = v mod 1000
  - If a page overflows, then use an extra overflow page
  - At 90% load on pages, 1.2 block accesses per request!
  - BUT, doesn't help for key range access ($10 < v < 75$)

B-Tree Structure

- Index node is a sequence of [pointer, key] pairs
- $K_1 < K_2 < \ldots < K_{n-2} < K_{n-1}$
- P_1 points to a node containing keys $< K_1$
- P_i points to a node containing keys in range $[K_{i-1}, K_i)$
- P_n points to a node containing keys $> K_{n-1}$
- So, $K'_1 < K'_2 < \ldots < K'_{n-2} < K'_{n-1}$



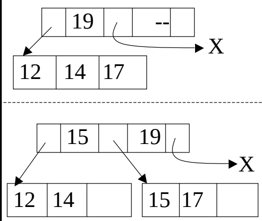
Example n=3



- Notice that leaves are sorted by key, left-to-right
- Search for value v by following path from the root
- If key = 8 bytes, ptr = 2 bytes, page = 4K, then n = 409
- So 3-level index has up to 68M leaves (409^3)
- At 20 records per leaf, that's 136M records

Insertion

- To insert key v, search for the leaf where v should appear
- If there's space on the leave, insert the record
- If no, split the leaf in half, and split the key range in its parent to point to the two leaves



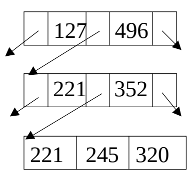
To insert key 15
- split the leaf
- split the parent's range [0, 19) to [0, 15) and [15, 19)
- if the parent was full, you'd split that too (not shown here)
- this automatically keeps the tree balanced

B-Tree Observations

- Delete algorithm merges adjacent nodes < 50% full, but rarely used in practice
- Root and most level-1 nodes are cached, to reduce disk accesses
- Secondary (non-clustered) index - Leaves contain [key, record id] pairs.
- Primary (clustered) index - Leaves contain records
- Use key prefix for long (string) key values
  - drop prefix and add to suffix as you move down the tree

Key Range Locks

- Lock on B-tree key range is a cheap predicate lock

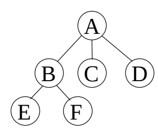


- Select Dept Where ((Budget > 250) and (Budget < 350))
- lock the key range [221, 352) record
- only useful when query is on an indexed field

- Commonly used with multi-granularity locking
  - Insert/delete locks record and intention-write locks range
  - MGL tree defines a fixed set of predicates, and thereby avoids predicate satisfiability

5.11 Tree Locking

- Can beat 2PL by exploiting root-to-leaf access in a tree
- If searching for a leaf, after setting a lock on a node, release the lock on its parent

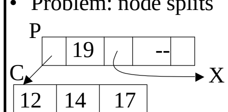


wl(A) wl(B) wu(A) wl(E) wu(B)

- The lock order on the root serializes access to other nodes

B-tree Locking

- Root lock on a B-tree is a bottleneck
- Use tree locking to relieve it
- Problem: node splits

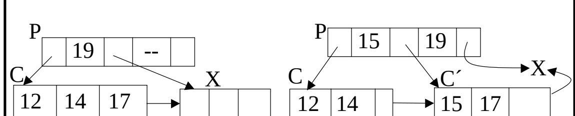


If you unlock P before splitting C, then you have to back up and lock P again, which breaks the tree locking protocol.

- So, don't unlock a node till you're sure its child won't split (i.e. has space for an insert)
- Implies different locking rules for different ops (search vs. insert/update)

B-link Optimization

- B-link tree - Each node has a side pointer to the next
- After searching a node, you can release its lock before locking its child
  - $r_1[P]\ r_2[P]\ r_2[C]\ w_2[C]\ w_2[C']\ w_2[P]\ r_1[C]\ r_1[C']$



- Searching has the same behavior as if it locked the child before releasing the parent … and ran later (after the insert)