CSE462/562: Database Systems (Fall 24)

Lecture 17: Transaction, Pessimistic Concurrency Control & Crash Recovery

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What is a transaction?

Transaction:

```
BEGIN;
INSERT INTO A VALUES (…)
SELECT * from A;
DELETE FROM A WHERE …;
COMMIT;
```

- A transaction is a sequence of one or more SQL operations treated as a unit
  - START/BEGIN [TRANSACTION] to start a new transaction
  - COMMIT: make all the changes by the current transaction permanent and visible
  - ROLLBACK/ABORT: revert all the changes by the current transaction

Recap on Transactions & Concurrency

- ACID properties
  - Atomicity
    - A Xact's effect is always applied as a whole, or not at all
  - Consistency
    - Run by itself must leave the DB in a consistent state (no IC violations)
  - Isolation
    - "protected" from the effects of concurrently scheduled other transactions
    - Most stringent isolation level: *serializable*
      - *Operations may be interleaved, but execution must be equivalent to some sequential (serial) order of all transactions*
  - Durability
    - If a transaction has successfully completed, its effects should persist even if the system crashes before all its changes are reflected on disk.

- Issues: Effect of interleaving transactions, and crashes, may result violate ACID.
  - Needs *concurrency control* & crash recovery

Scheduling Transactions

- *Serial schedule:* Schedule that does not interleave the actions of different transactions.
- *Equivalent schedules*: For any database state, the effect of executing the first schedule is identical to the effect of executing the second schedule.
- *Serializable schedule*: A schedule that is equivalent to some serial execution of the transactions.

  (Note: If each transaction preserves consistency, every serializable schedule preserves consistency.)

- When we discuss schedules, we only consider reads/writes/commit/abort
  - Ignores computation
- Two forms of (restricted) serializability
  - conflict serializable
  - view serializability

Anomalies with interleaved execution

- Dirty reads (WR conflict)

T1:	R(A), W(A),		R(B), W(B), Abort
T2:		R(A), W(A), C	

- Unrepeatable reads (RW conflict)

T1:	R(A),		R(A), W(A), C
T2:		R(A), W(A), C	

Anomalies with interleaved execution

- Phantom read (RW conflict w/ predicate)

T1:	R(t: P(t))		R(t: P(t)) C
T2:		W(A' , s.t. $A' \in \boldsymbol{P}$) C	

- Dirty write (WW conflict)

T1:	W(A)		W(B) C
T2:		W(A) W(B) C	

Conflict serializability

- Two operations of two different transactions conflict if
  - Performed on the same object
  - At least one of them is a write

T1:	R$_1$ (A), W_1(A),		R_1(B), W_1(B)
T2:		R_2(A), W_2(A)	

Conflicts:

$R_1(A), W_2(A)$
$W_1(A), R_2(A)$
$W_1(A), W_2(A)$

- We can swap two adjacent nonconflicting operations without changing the final state

T1:	R$_1$ (A), W_1(A), R_1(B), W_1(B)	
T2:		R_2(A), W_2(A)

- Two schedules are conflict equivalent if one can be transformed into the other through swaps
  - Involve the same actions of the same transactions in the same order
  - Every pair of conflicting operations are ordered the same way
- Schedule S is said to be conflict serializable if it is *conflict equivalent* to some *serial* schedule S'

View serializability

- View serializability is based on view equivalence
- Schedules S1 and S2 are view equivalent if:
  - If Ti reads initial value of A in S1, then Ti also reads initial value of A in S2
  - If Ti reads value of A written by Tj in S1, then Ti also reads value of A written by Tj in S2
  - If Ti writes final value of A in S1, then Ti also writes final value of A in S2

T1:	R(A)		W(A)	
T2:		W(A)		
T3:				W(A)

T1:	R(A),W(A)		
T2:		W(A)	
T3:			W(A)

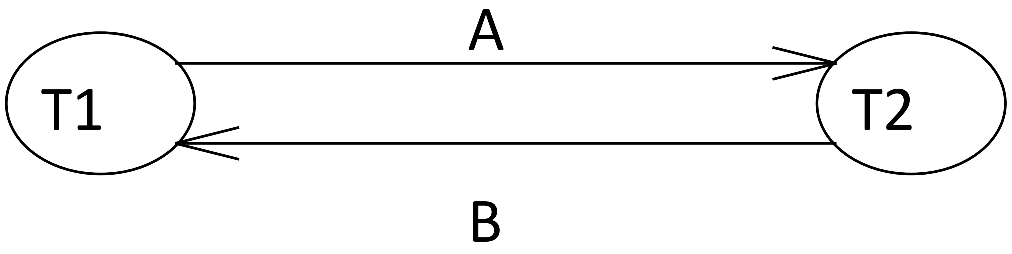
View equivalent but not conflict equivalent

- View serializability is "weaker" than conflict serializability!
  - Every conflict serializable schedule is view serializable, but not vice versa!
  - I.e. admits more serializable schedules

Determining conflict serializability

- Dependency graph
  - One node per Xact
    - edge from *Ti* to *Tj* if
      - an operation of Ti conflicts with an operation of Tj and
      - Ti’s operation appears earlier in the schedule than the conflicting operation of Tj.
- <u>Theorem</u>: Schedule is conflict serializable if and only if its dependency graph is acyclic

T1:	R(A), W(A),		R(B), W(B)
T2:		R(A), W(A), R(B), W(B)	



Dependency graph

How to enforce conflict serializability?

- Two operations of two different transactions conflict if
  - Performed on the same object
  - At least one of them is a write

T1:	R_1 (A), W_1(A),		R_1(B), W_1(B)
T2:		R_2(A), W_2(A)	

Conflicts:
$R_1(A), W_2(A)$
$W_1(A), R_2(A)$
$W_1(A), W_2(A)$

- We can swap two adjacent nonconflicting operations without changing the final state

T1:	R_1 (A), W_1(A), R_1(B), W_1(B)	
T2:		R_2(A), W_2(A)

- Two schedules are conflict equivalent if one can be transformed into the other through swaps
  - Involve the same actions of the same transactions in the same order
  - Every pair of conflicting operations are ordered the same way
- Schedule S is said to be conflict serializable if it is *conflict equivalent* to some *serial* schedule S'

Pessimistic Concurrency Control

- *Strict Two-phase Locking (Strict 2PL) Protocol*:
  - Each Xact must obtain a S (*shared*) lock on object before reading, and an X (*exclusive*) lock on object before writing.
  - All locks held by a transaction are released when the transaction completes
    - (Non-strict) 2PL Variant: Release locks anytime, but cannot acquire locks after releasing any lock.
  - If an Xact holds an X lock on an object, no other Xact can get a lock (S or X) on that object.
- Strict 2PL allows only conflict serializable schedules.
  - Additionally, it simplifies transaction aborts
  - (Non-strict) 2PL also allows only serializable schedules, but involves more complex abort processing

Lock Compatibility Matrix

	S	**X**
S	√	–
X	–	–

Example: strict 2-PL

A

B

T1: A = A + 100, B = B - 100
T2: A = A - 100, B = B + 100

Lock upgrade

T1	T2
S(A)	
R(A)	
X(A)	
W(A)	
	request S(A) -- blocked
S(B)	
R(B)	
X(B)	
W(B)	
Commit	
Release A & B	
	S(A)
	R(A)
	X(A)
	W(A)
	……

Example: non-strict 2-PL

A

B

T1: A = A + 100, B = B - 100
T2: A = A - 100, B = B + 100

No new locks/lock upgrades at this point.

T1	T2
X(A)	
X(B)	
R(A)	
W(A)	
	request S(A) -- blocked
Release A	
	S(A)
	R(A)
	X(A)
	W(A)
R(B)	
W(B)	
Release B	
Commit	
	……

Example: non-strict 2-PL

A

B

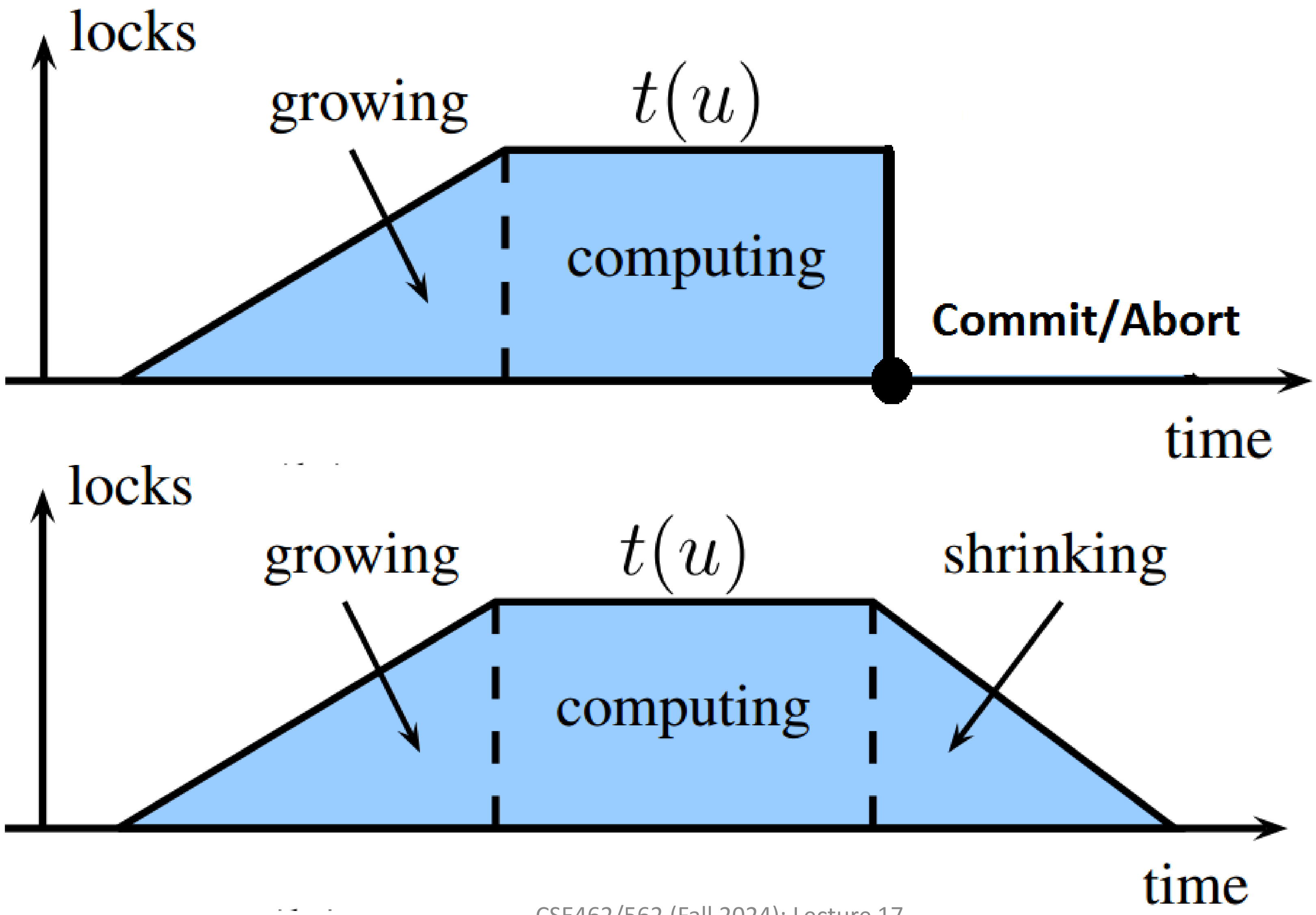
T1: A = A + 100, B = B - 100
T2: A = A - 100, B = B + 100

susceptible to cascading aborts!

Usually avoided in DBMS to avoid wasted work.

T1	T2
X(A)	
X(B)	
R(A)	
W(A)	
	request S(A) -- blocked
Release A	
	S(A)
	R(A)
	X(A)
	W(A)
R(B)	
W(B)	
Release B	
abort	
	abort

Strict 2-PL vs non-strict 2-PL



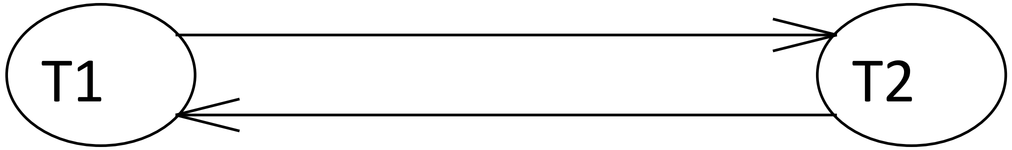
Deadlocks

A

B

T1: A = A + 100, B = B - 100
T2: B = B + 100, A = A - 100

- Create a waits-for graph:
  - Nodes are transactions
  - There is an edge from Ti to Tj if Ti is waiting for Tj to release a lock
- Deadline ⇔ cycle in the wait-for graph
- Two ways to handle deadlocks
  - Deadlock prevention
  - Deadlock detection



T1	T2
S(A)	
R(A)	
X(A)	
W(A)	
	S(B)
	R(B)
	X(B)
	W(B)
S(B) -- blocked	
	S(A) -- blocked

Deadlock!

Deadlock prevention

- Idea: make sure wait-for graph is acyclic
  - Intuition: only allow edges to form in one of the following two directions:
    - either from older transactions to younger transactions *(wait-die)*
    - or only from younger to older *(wound-wait)*
  - Aborting a transaction prevents forming wait-for edges

- Assign priorities based on start timestamps.
  Assume Ti wants a lock that Tj holds. Two policies are possible:
  - *Wait-Die:* If Ti has lower timestamp (i.e., older) than Tj, Ti waits; otherwise Ti aborts
    - No preemption
  - *Wound-Wait:* If Ti has lower timestamp (i.e., older), Tj aborts (preempted); otherwise Ti waits
    - Preemptive scheduling

- If a transaction re-starts, make sure it gets its original timestamp
  - Why? (to avoid starvation)

Deadlock prevention: Wait-Die

A

B

T1: A = A + 100, B = B - 100
T2: B = B + 100, A = A - 100

Wait-Die: If Ti has lower timestamp (i.e., older) than Tj, Ti waits; otherwise Ti aborts

Scenario 1: T1 requests $S(B)$ before $T2$ requests $S(A)$

T1, ts = 1	T2, ts = 2
S(A)	
R(A)	
X(A)	
W(A)	
	S(B)
	R(B)
	X(B)
	W(B)
S(B) -- blocked	
	S(A) -- abort
S(B) granted	
R(B)	
X(B)	
W(B)	
commit	
	(retry with ts = 2…)

Deadlock prevention: Wait-Die

A

B

T1: A = A + 100, B = B - 100
T2: B = B + 100, A = A - 100

Wait-Die: If Ti has lower timestamp (i.e., older) than Tj, Ti waits; otherwise Ti aborts

Scenario 2: T1 requests $S(B)$ after $T2$ requests $S(A)$

T1, ts = 1	T2, ts = 2
S(A)	
R(A)	
X(A)	
W(A)	
	S(B)
	R(B)
	X(B)
	W(B)
	S(A) -- abort
S(B) granted	
R(B)	
X(B)	
W(B)	
commit	
	(retry with ts = 2…)

T1

T2

Deadlock prevention: Wound-Wait

A

B

T1: A = A + 100, B = B - 100
T2: B = B + 100, A = A - 100

Wound-Wait: If Ti has lower timestamp (i.e., older), Tj aborts (preempted); otherwise Ti waits

Scenario 1: T1 requests $S(B)$ before $T2$ requests $S(A)$

T1

T2

T1, ts = 1	T2, ts = 2
S(A)	
R(A)	
X(A)	
W(A)	
	S(B)
	R(B)
	X(B)
	W(B)
S(B)	abort *(preempted)*
R(B)	
X(B)	
W(B)	
commit	
	(retry with ts = 2…)

Deadlock prevention: Wound-Wait

A

B

T1: A = A + 100, B = B - 100
T2: B = B + 100, A = A - 100

Wound-Wait: If Ti has lower timestamp (i.e., older), Tj aborts (preempted); otherwise Ti waits

Scenario 2: T1 requests $S(B)$ after $T2$ requests $S(A)$

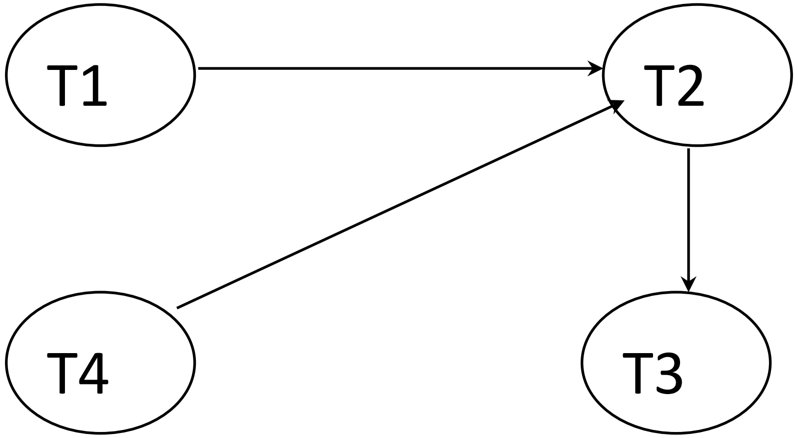
T1, ts = 1	T2, ts = 2
S(A)	
R(A)	
X(A)	
W(A)	
	S(B)
	R(B)
	X(B)
	W(B)
	S(A) -- blocked
S(B)	abort *(preempted)*
R(B)	
X(B)	
W(B)	
commit	
	(retry with ts = 2…)

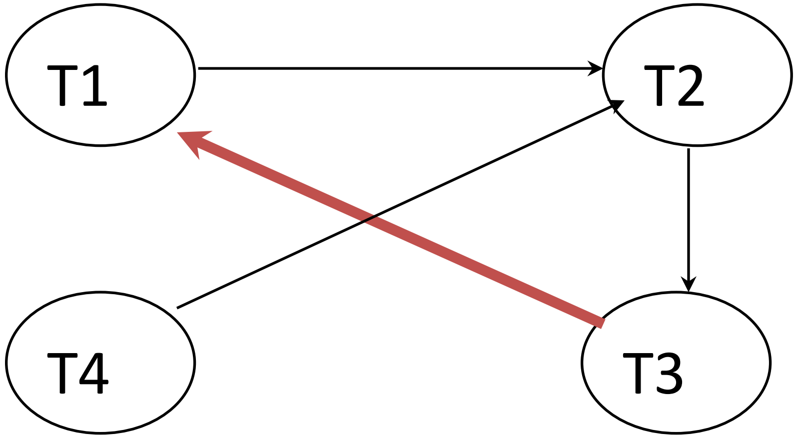
wait-for edge from T2 to T1 disappears after T2 is preempted

Deadlock detection

- Explicitly create a waits-for graph:
  - Nodes are transactions
  - There is an edge from Ti to Tj if Ti is waiting for Tj to release a lock
- Periodically check for cycles in the waits-for graph
  - If there's a cycle, abort at least one transaction in the cycle

T1: S(A), S(D), S(B)
T2: X(B) X(C)
T3: S(D), S(C), X(A)
T4: X(B)





Deadlock detection (cont'd)

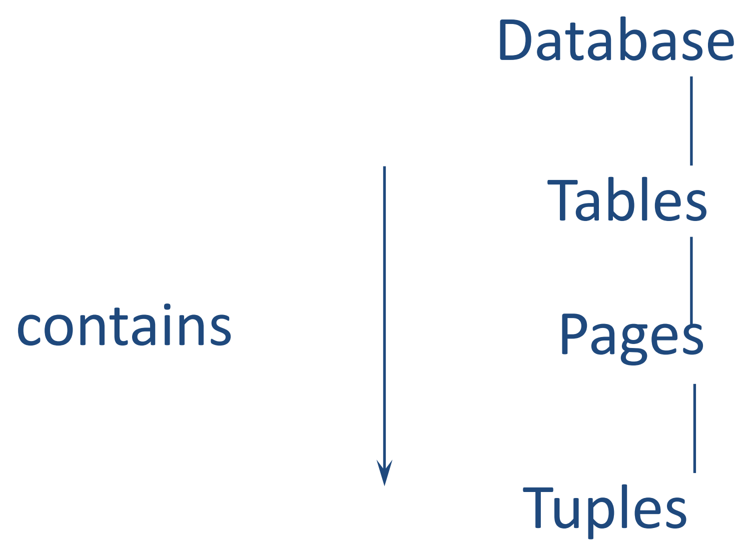
- In practice, most systems do detection
  - Experiments show that most waits-for cycles are length 2 or 3
  - Hence, only a few transactions actually need to be aborted
  - Implementations can vary
    - Can construct the graph and periodically look for cycles
      - When is the graph created ?
      - Which process checks for cycles ?
    - Can also use a "time-out" scheme
      - if T has been waiting on a lock for a long time, assume it's in a deadlock and abort

What we have glossed over

- What should we lock?
  - We assume tuples here, but that can be expensive!
  - If we do table locks, that’s too conservative
  - *Multi-granularity* locking
- How to deal with phantoms?
- Locking in indexes
  - don’t want to lock a B-tree root for a whole transaction!
  - more fine-grained concurrency control in indexes
- CC w/out locking (we’ll omit it in this course)
  - “optimistic” concurrency control
  - “timestamp” and multi-version concurrency control
  - locking usually better, though

Multi-granularity locks

- Hard to decide what granularity to lock (tuples vs. pages vs. tables).
- Shouldn't have to make same decision for all transactions!
- Data "containers" are nested:



Solution: new lock modes and protocols

- Allow Xacts to lock at each level, but with a special protocol using new "intention" locks:
- Still need S and X locks, but before locking an item, Xact must have proper intension locks on all its ancestors in the granularity hierarchy.

- IS – Intent to get S lock(s) at finer granularity.
- IX – Intent to get X lock(s) at finer granularity.
- SIX mode: Like S & IX at the same time. Why useful?

	IS	IX	SIX	S	X
IS	√	√	√	√	
IX	√	√			
SIX	√				
S	√			√	
X					

Example: 2-level hierarchy

- T1 scans R, and updates a few tuples:
  - T1 gets an SIX lock on R, then get X lock on tuples that are updated.
- T2 uses an index to read only part of R:
  - T2 gets an IS lock on R, and repeatedly gets an S lock on tuples of R.
- T3 reads all of R:
  - T3 gets an S lock on R.
  - OR, T3 could behave like T2; can use lock escalation to decide which.
- Lock escalation
  - Dynamically asks for coarser-grained locks when too many low level locks acquired

Tables

Tuples

	IS	IX	SIX	S	X
IS	√	√	√	√	
IX	√	√			
SIX	√				
S	√			√	
X					

Dynamic Databases – The "Phantom" Problem

- If the DB is not a fixed collection of objects, even Strict 2PL (on individual items) will not assure serializability:
- Consider T1 – "Find the highest GPA among students of each age"
  - T1 locks all pages containing sailor records with *age* = 20
    - and finds the highest GPA (say, *GPA* = 3.7).
  - Next, T2 inserts a new student; *GPA* = 4.0, *age* = 20.
  - T2 also deletes student with the highest GPA (say 3.8) among those of age = 21, and commits.
  - T1 now locks all pages containing student records with age = 21, and finds highest GPA (say, *GPA* = 3.6).
- No serial execution could lead to T1's result!

The problem

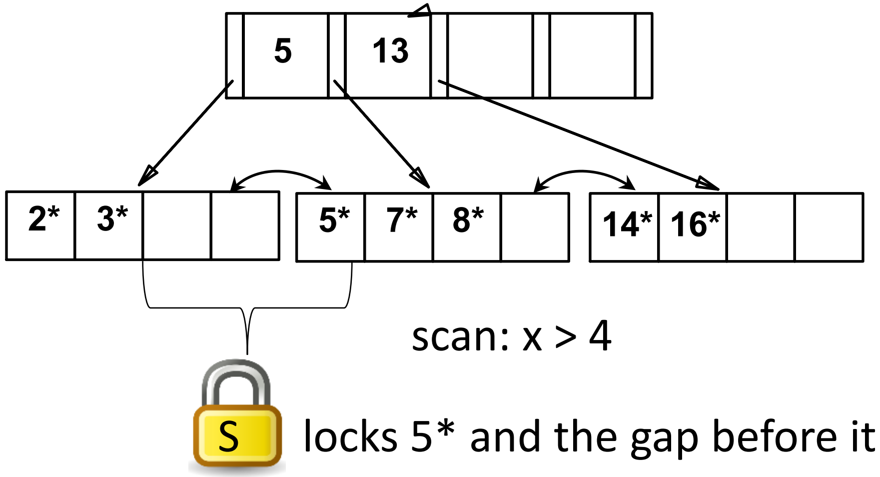
- T1 implicitly assumes that it has locked the set of all student records with *age* = 20.
  - Assumption only holds if no student records are added while T1 is executing!
  - Need some mechanism to enforce this assumption. (Index locking and predicate locking.)
- Example shows that conflict serializability guarantees serializability only if the set of objects is fixed!
  - e.g. table locks
- Solution: predicate locking

Predicate locking

- Grant lock on all records that satisfy some logical predicate, e.g. *age > 2*salary*.
- Index locking is a special case of predicate locking for which an index supports efficient implementation of the predicate lock.
  - What is the predicate in the sailor example?
- General predicate locking has a lot of locking overhead.
  - too expensive!

Instead of predicate locking

- Full table scans lock entire tables
- Range lookups do “next-key” & gap locking
  - physical stand-in for a logical range!



scan: x > 4

locks 5* and the gap before it (3, 5)

At this point,

insert 4: blocked

insert 10?

Lock management

- Lock and unlock requests are handled by the lock manager
- Lock table: a hash table over lock table entries
  - for various resources, e.g., records, gaps, pages, tables, …
- Lock table entry:
  - Number of transactions currently holding a lock
  - Type of lock held (S, X, IS, IX, SIX)
  - Pointer to queue of lock requests
- Locking and unlocking have to be atomic operations
  - requires *latches* (e.g. reader-writer locks/semaphores), which ensure that the process is not interrupted while managing lock table entries
- Lock upgrade: transaction that holds a shared lock can be upgraded to hold an exclusive lock
  - Can cause deadlock problems
- Deadlock prevention/detection

Locks vs Latches

- What’s common ?
  - Both used to synchronize concurrent tasks
- What’s different ?
  - Locks are used for *logical consistency*
  - Latches are used for *physical consistency*
- Why treat ‘em differently ?
  - Latches are short-duration lower-level locks that protects critical sections in the code
    - depends on DBMS developer to prevent deadlocks
  - Locks protects data/resources, much longer duration
    - need deadlock prevention/detection, aborting transactions using priorities
    - more lock modes, hierarchical
- Where are latches used ?
  - In a lock manager !
  - In a shared memory buffer manager
  - In a B+ Tree index
  - In a log/transaction/recovery manager

Locks vs Latches

	Latches	**Locks**
Ownership	Processes	Transactions
Duration	Very short	Long (Xact duration)
Deadlocks	No detection - code carefully !	Checked for deadlocks
Overhead	Cheap - 10s of instructions (latch is directly addressable)	Costly - 100s of instructions (have to search for lock)
Modes	S, X	S, X, IS, IX, SIX
Granularity	Flat - no hierarchy	Hierarchical

Recap on Transactions & Concurrency

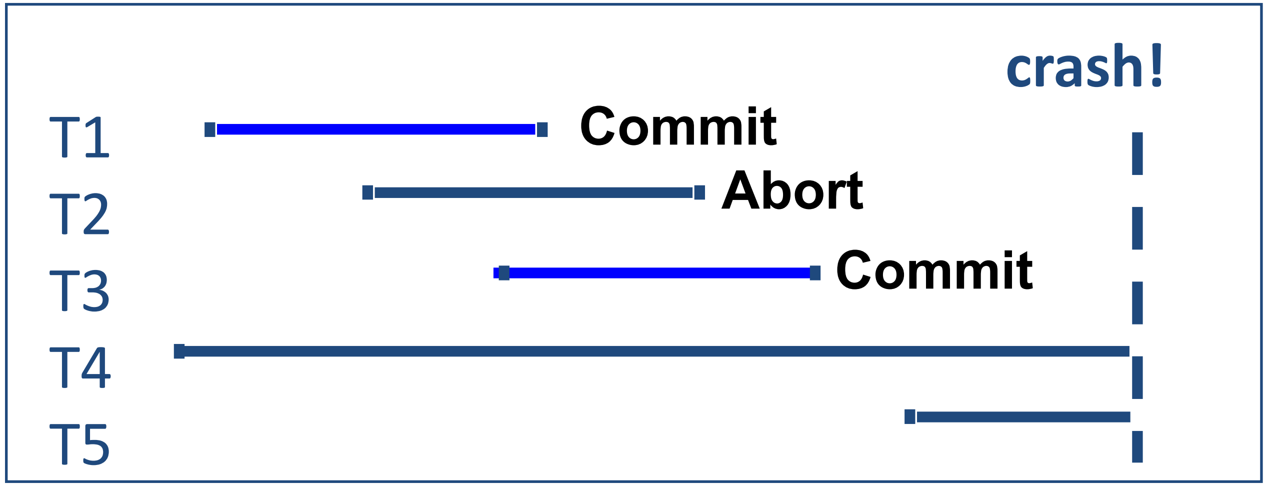
- Atomicity
  - A Xact's effect is always applied as a whole, or not at all
- Consistency
  - Run by itself must leave the DB in a consistent state (no IC violations)
- Isolation
  - "protected" from the effects of concurrently scheduled other transactions
- Durability
  - If a transaction has successfully completed, its effects should persist even if the system crashes before all its changes are reflected on disk.

- Issues: Effect of interleaving transactions, and crashes, may result violate ACID.
  - Needs concurrency control & *crash recovery*

Motivation for crash recovery

- Atomicity:
  - Transactions may abort (“Rollback”).
- Durability:
  - What if DBMS stops running? (Causes?)

- Desired state after system restarts:
  - T1 & T3 should be durable.
  - T2, T4 & T5 should be aborted (effects not seen).



Assumptions

- Concurrency control is in effect.
  - Strict 2-PL, in particular.
- Updates are happening "in place".
  - i.e. data are overwritten on (or deleted from) the actual pages.

- Can you think of a *simple* scheme (requiring no logging) to guarantee Atomicity & Durability?
  - What happens during normal execution (what is the minimum lock granularity)?
  - What happens when a transaction commits?
  - What happens when a transaction aborts?

Buffer manager plays a key role

- **Force policy –** make sure that every update is on disk before commit.
  - Provides durability without REDO logging.
  - But, can cause poor performance.

- **No Steal policy –** don't allow buffer-pool frames with *uncommited* updates to overwrite *committed* data on disk.
  - Useful for ensuring atomicity without UNDO logging.
  - But can cause poor performance.

Preferred buffer management policy: steal/no-force

- This combination is most complicated but allows for highest performance.
- *NO FORCE*: do not have to flush all dirty pages of a transaction to disk before it commits
  - complicates Durability
  - What if system crashes before a modified page written by a committed transaction makes it to disk?
  - Write as little as possible, in a convenient place, at commit time, to support REDOing modifications.
- *STEAL:* allows buffer pool with uncommitted updates to overwrite committed data on disk
  - complicates Atomicity
  - What if the Xact that performed updates aborts?
  - What if system crashes before Xact is finished?
  - Must remember the old value of P (to support UNDOing the write to page P).

Buffer management policies

	No Steal	Steal
No Force		Fastest
Force	Slowest	

Performance Implications

	No Steal	Steal
No Force	No UNDO REDO	UNDO REDO
Force	No UNDO No REDO	UNDO No REDO

Logging/Recovery Implications

Basic Idea: Logging

- Record REDO and UNDO information, for every update, in a *log*.
  - Sequential writes to log (put it on a separate disk).
  - Minimal info (diff) written to log, so multiple updates fit in a single log page.
- Log: An ordered list of REDO/UNDO actions
  - Log record contains:

    <XID, pageID, offset, length, old data, new data>
  - and additional control info (which we'll see soon).

Write-Ahead Logging (WAL)

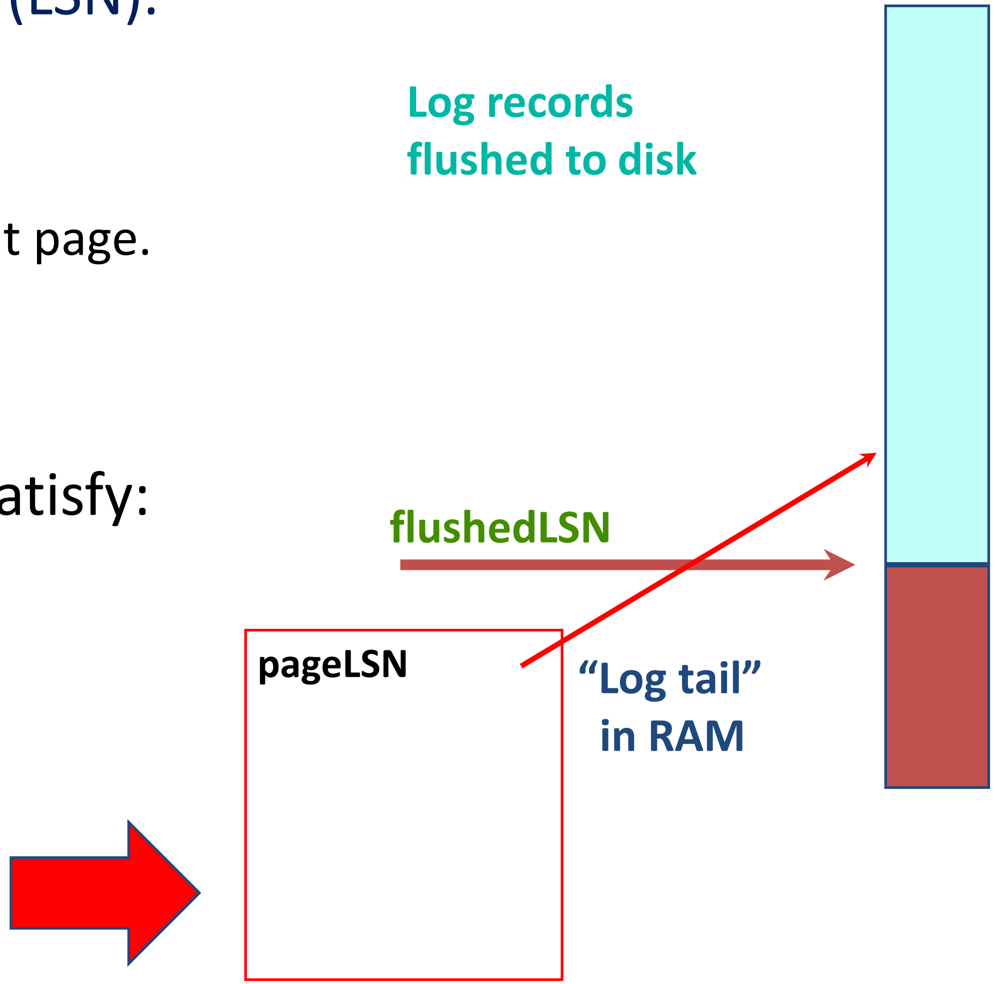
- The Write-Ahead Logging Protocol:
  - ① Must flush the log record for an update *before* the corresponding data page gets to disk.
  - ② Must flush all log records for a Xact *before commit*
    - alternatively,. transaction is not considered as committed until all of its log records including its “commit” record are on the stable log.
- #1 (with UNDO info) helps provide Atomicity.
- #2 (with REDO info) helps provide Durability.
- This allows us to employ Steal/No-Force policy

- Exactly how is logging (and recovery) done?
  - We’ll look at the ARIES algorithms.
    - Algorithms for Recovery and Isolation Exploiting Semantics

WAL & the log

- Each log record has a unique Log Sequence Number (LSN).
  - LSNs are monotonically increasing.
- Each *data page* contains a pageLSN.
  - The LSN of the *most recent log record* for an update to that page.
- System keeps track of flushedLSN.
  - The max LSN flushed so far.
- WAL: Before page i is flushed to disk, the log must satisfy:

$$pageLSN_i \leq flushedLSN$$



Log Records

LogRecord fields:

LSN
prevLSN
XID
type
pageID
length
offset
before-image
after-image

update records only (pageID, length, offset, before-image, after-image)

prevLSN is the LSN of the previous log record written by *this* Xact (so records of an Xact form a linked list backwards in time)

Possible log record types:

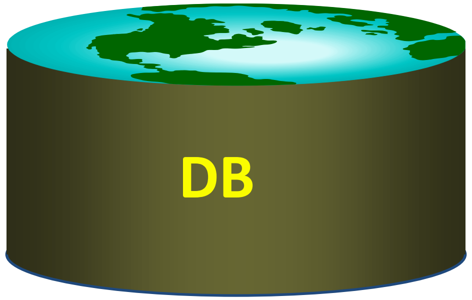
- Update
- Checkpoint (for log maintenance)
- Compensation Log Records (CLRs)
  - for UNDO actions
- Commit/Abort
- End (indicates end of commit/abort)

Other logging-related state

- Two -in-memory tables
- Transaction Table
  - One entry per currently active Xact.
    - entry removed when Xact commits or aborts
  - Contains XID, status (running/committing/aborting), and lastLSN (most recent LSN written by Xact).
- Dirty Page Table:
  - One entry per dirty page currently in buffer pool.
  - Contains recLSN -- the LSN of the log record which ***first*** caused the page to be dirty.
    - If a dirty page is flushed to disk, it is removed from dirty page table

The big picture: what's stored and where

LogRecords
- LSN
- prevLSN
- XID
- type
- pageID
- length
- offset
- before-image
- after-image



Data pages
each
with a
pageLSN

Master record

Xact Table
- lastLSN
- status

Dirty Page Table
- recLSN

flushedLSN

Normal execution of an Xact

- Series of reads & writes, followed by commit or abort.
  - We will assume that disk write is atomic.
    - In practice, additional details to deal with non-atomic writes.
- Strict 2-PL.
- STEAL, NO-FORCE buffer management, with Write-Ahead Logging.

Transaction Commit

- Write commit record to log.
- All log records up to Xact's commit record are flushed to disk.
  - Guarantees that flushedLSN $\geq$ lastLSN.
  - Note that log flushes are sequential, synchronous writes to disk.
  - Many log records per log page.
- Write an end record to log (no need to flush immediately)
- Commit() returns.

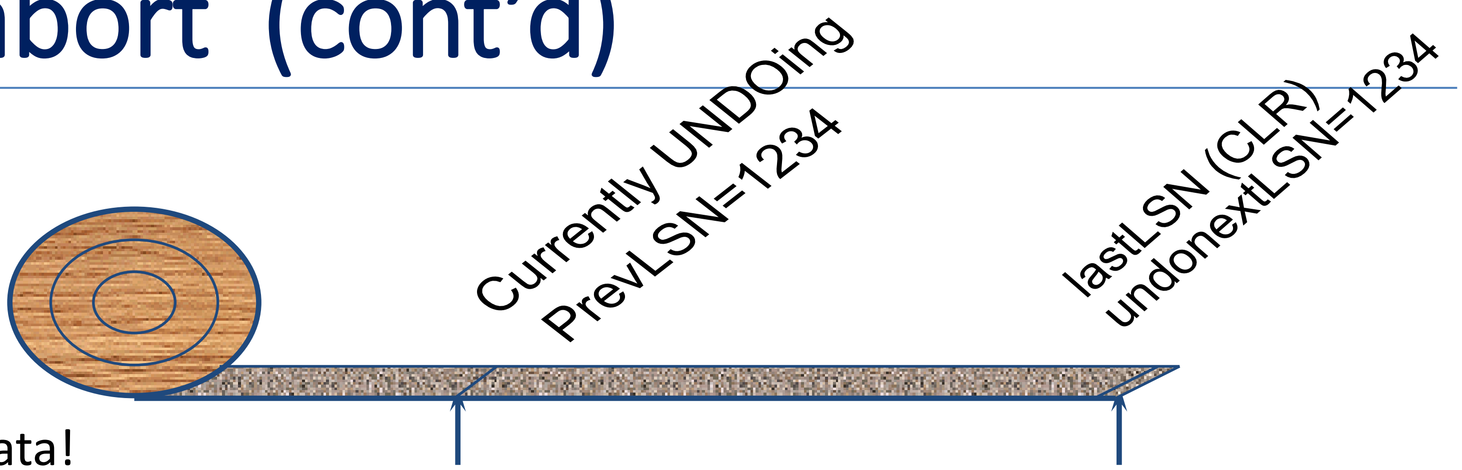
- When does a transaction becomes durable in the database?
  - When its commit log record is flushed to disk, even if there are still dirty pages in bufmgr.

Simple transaction abort

- For now, consider an explicit abort of a Xact.
  - No crash involved.
- First, set the transaction state in the transaction table to aborting.
  - Write an *Abort* log record before starting to rollback operations
- We want to "play back" the log in reverse order, UNDOing updates.
  - Get lastLSN of Xact from Xact table.
    - Can follow chain of log records backward via the prevLSN field.
  - Write a "CLR" (compensation log record) for each undone operation.
    - more details on next slide
  - Once its finished, write a transaction end log record in the disk

- Q: do we need to wait for abort, CLRs and end record to be flushed?

Simple transaction abort (cont'd)

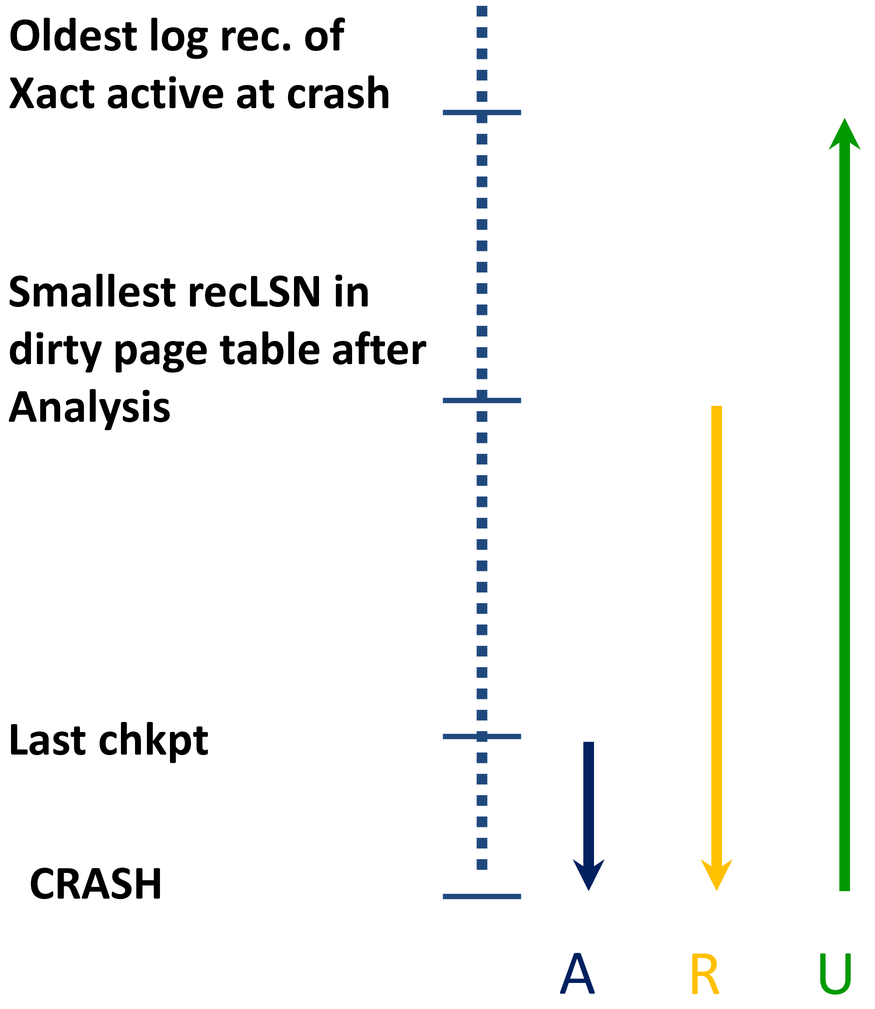


- To perform UNDO, must have a lock on data!
  - We still have the lock because of strict 2-PL.
- Before restoring old value of a page, write a CLR:
  - Must continue logging during undo in case of crash
  - CLR has one extra field: undonextLSN
    - Points to the next LSN to undo (i.e. the prevLSN of the record we're currently undoing).
  - CLR contains REDO info
  - CLRs is *never* undone
    - Undo needn't be idempotent (>1 UNDO won't happen)
    - But they might be Redone when repeating history (=1 UNDO guaranteed)
- At end of all UNDOs, write an "end" log record.

Checkpointing

- Conceptually, we keep log around for all time. Obviously this has performance issues…
- Periodically, the DBMS creates a checkpoint, in order to minimize the time taken to recover in the event of a system crash. Write to log:
  - begin_checkpoint record: Indicates when chkpt began.
  - end_checkpoint record: Contains current *Xact table* and *dirty page table*. This is a `fuzzy checkpoint':
    - Other Xacts continue to run; so these tables accurate only as of the time of the begin_checkpoint record.
    - No attempt to force all dirty pages to disk; effectiveness of checkpoint limited by oldest unwritten change to a dirty page.
      - However, the more dirty page gets flushed, the shorter time will be needed in crash recovery
  - Store LSN of most recent chkpt record in a safe place (*master* record).

Crash Recovery: Big Picture



- Start from a checkpoint (found via master record).
- Three phases. Need to do:
  - Analysis - Figure out which Xacts committed since checkpoint, which failed.
  - REDO ***all*** actions.

    (repeat history)
  - UNDO effects of failed Xacts.

Phase 1: the analysis phase

- Re-establish knowledge of state at checkpoint.
  - via transaction table and dirty page table stored in the checkpoint
- Scan log forward from checkpoint.
  - End record: Remove Xact from Xact table.
  - All Other records: Add Xact to Xact table, set lastLSN=LSN, change Xact status on commit.
  - also, for Update records: If page P not in Dirty Page Table, Add P to DPT, set its recLSN=LSN.
- At end of Analysis…
  - transaction table says which xacts were active at time of crash.
  - DPT says which dirty pages *might not* have made it to disk

Phase 2: the redo phase

- We *Repeat History* to reconstruct state at crash:
  - Reapply *all* updates (including those of aborted Xacts), redo CLRs.
- Scan forward from log rec containing smallest recLSN in DPT.   Q: why start here?
- For each update log record or CLR with a given LSN, REDO the action unless:
  - Affected page is not in the Dirty Page Table, or
  - Affected page is in D.P.T., but has recLSN > LSN, or
  - pageLSN (in DB) $\geq$ LSN. (this last case requires I/O)
- To REDO an action:
  - Reapply logged action.
  - Set pageLSN to LSN. No additional logging, no forcing!

Phase 3: the undo phase

ToUndo={lastLSNs of all Xacts in the Trans Table}

i.e., last log entry of the aborted transactions

Repeat:

- Choose (and remove) largest LSN among ToUndo.
- If this LSN is a CLR and undonextLSN==NULL
  - Write an End record for this Xact.
- If this LSN is a CLR, and undonextLSN != NULL
  - Add undonextLSN to ToUndo
- Else this LSN is an update. Undo the update, write a CLR, add prevLSN to ToUndo.

Until ToUndo is empty.

Example of recovery

RAM

Xact Table

lastLSN

status

Dirty Page Table

recLSN

flushedLSN

ToUndo

LSN	LOG
00	begin_checkpoint
05	end_checkpoint
10	update: T1 writes P5
20	update T2 writes P3
30	T1 abort
40	CLR: Undo T1 LSN 10
45	T1 End
50	update: T3 writes P1
60	update: T2 writes P5
	CRASH, RESTART

prevLSNs

Example: crash during recovery

RAM

Xact Table
- lastLSN
- status

Dirty Page Table
- recLSN

flushedLSN

ToUndo

LSN	LOG
00,05	begin_checkpoint, end_checkpoint
10	update: T1 writes P5
20	update T2 writes P3
30	T1 abort
40	CLR: Undo T1 LSN 10
45	T1 End
50	update: T3 writes P1
60	update: T2 writes P5
	CRASH, RESTART
70	CLR: Undo T2 LSN 60
80	CLR: Undo T3 LSN 50
85	T3 end
	CRASH, RESTART
90,95	CLR: Undo T2 LSN 20, T2 end

undonextLSN

Additional crash issues

- What happens if system crashes during Analysis? During REDO?
- How do you limit the amount of work in REDO?
  - Flush asynchronously in the background.
  - Watch “hot spots”!
- How do you limit the amount of work in UNDO?
  - Avoid long-running Xacts.

- What about schema changes/disk space management?

Summary of logging/recovery

- Recovery Manager guarantees Atomicity & Durability.
- Use WAL to allow STEAL/NO-FORCE w/o sacrificing correctness.
- LSNs identify log records; linked into backwards chains per transaction (via prevLSN).
- pageLSN allows comparison of data page and log records.

- Checkpointing: A quick way to limit the amount of log to scan on recovery.
- Recovery works in 3 phases:
  - Analysis: Forward from checkpoint.
  - Redo: Forward from oldest recLSN.
  - Undo: Backward from end to first LSN of oldest Xact alive at crash.
- Upon Undo, write CLRs.
- Redo "repeats history": Simplifies the logic!