An Adaptable Constrained Locking Protocol for High Data Contention Environments

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Multiversion concurrency control schemes are often limited in their practicability due to their storage requirements for multiple versions of the data. However, a class of multiversion schemes utilize only the versions, maintained for the purpose of recovery, to improve the concurrency by allowing the concurrent execution of "non conflicting" read-write lock requests on different versions of data in an arbitrary fashion. A transaction that accesses a data item version which is later diagnosed to lead to an incorrect execution, is aborted. This act is reminiscent of the validation phase in the optimistic concurrency schemes. Various performance studies suggest that these schemes perform poorly in high data contention environments where the excessive transaction aborts result, due to the failed validation. We propose an adaptable constrained two version two phase locking (C2V2PL) scheme in which these "non conflicting" requests are allowed only in a constrained manner. C2V2PL scheme assumes that a lock request failing to satisfy the specific constraints will lead to an incorrect execution and hence, must be either rejected or blocked. This eliminates the need for a separate validation phase. When the contention for data among the concurrent transactions is high, the C2V2PL scheduler in aggressive state rejects such lock requests. The deadlock free nature of C2V2PL scheduler in this state further reduces the duration for which locks are held by a transaction. The C2V2PL scheduler adapts to the low data contention environments by accepting the lock requests that have failed the specific constraints but contrary to the assumption will not lead to an incorrect execution. Thus improving the performance due to reduced transaction aborts in this conservative state.
1 Introduction

Many multiversion concurrency control schemes using a bounded number of versions for the data items have been proposed for improving the performance of transaction processing. These schemes have been broadly categorized under mixed and pure multiversion schemes in [BHГ87]. The mixed multiversion schemes [CFL+82, Wei87, AS89, BC91] have two types of transactions, i.e. the read-only transactions and the update transactions. The read-only transactions read the old but consistent versions while the update transactions manipulate only the “current” version via two phase locking (2PL) protocol. Even if we assume that the transaction type can be determined for every transaction when it starts executing, which is not the case for at least the on-line transactions, the increase in the size and frequency of the update transactions because of increased acceptance of the transaction as an organizational concept for a wider variety of applications (e.g. the database servers [Val93] on the information superhighways), limits the performance of the system if only the “current” version is available for their synchronization. In high data contention applications like stock exchange databases [PR88], the mixed schemes will pose the problems for the update transactions same as in ordinary two phase locking schemes [TGS85].

Pure multiversion schemes using two phase locking [BHR80, SR81, BHГ87, KSI91] utilize the versions, maintained by the system for the reasons of recovery, for allowing the concurrent execution of the conflicting transactions. The two phase locking for write-write synchronization puts an upper bound on the number of versions for every data item. Since the concurrent access to the conflicting read-write actions is allowed on different versions of a data item in an unrestricted fashion, the execution of each transaction must be validated before its effects can be committed. This validation is usually performed at the end of the transaction execution, either because it is computationally expensive to validate each
action executed on behalf of the transaction [BHR80, SR81] or because the scheme does not allow any other validation point [BH87]. In any case, the effort in executing the transaction that fails the validation is wasted. These pure multiversion schemes will be recognized as optimistic concurrency schemes in the taxonomy of schedulers by [BH87]. In the optimistic schemes, the transaction aborts due to the failed validation grows rapidly with the increase in contention for data [ACL87]. The effect of these aborts on the system performance becomes more prominent as the size of the transaction grows.

In this paper, we present an adaptable Constrained Two Version Two Phase Locking (C2V2PL) scheme for synchronising the read and write lock requests on the different versions of a data item in only a constrained manner. The constraints are specified in terms of timestamps on the lock requested and on the locks held for the data item. The correctness of the transaction execution is guaranteed if the transaction can announce its completion, by submitting its commit action, to the scheduler. No separate validation phase for validating the transaction execution is required. A maximum of two committed versions of a data item are available at any given time. A read request is completed by using the Read rule similar to the multiversion timestamp ordering (MVTO) read rule in [BG83]. The action taken by the scheduler on the lock request that fails to satisfy the constraints is dependent on the scheduler state. When the conflicts for data is high, such lock requests are rejected and the scheduler is said to be in aggressive state. When the data contention is low, these lock requests are blocked and the scheduler is said to be in conservative state. In the aggressive state, since no lock request gets blocked for indefinite periods of time, the conflicting transactions never deadlock on a lock request. In the conservative state, the blocking of these lock requests may lead to deadlock, but may also improve the transaction throughput by avoiding the unnecessary abort of the transactions.
The rest of the paper is organized as follows. In section 2, we present the transaction model and the database model used in C2V2PL. We present the adaptable C2V2PL scheme in conservative and aggressive states in section 3. The comparative behavior of C2V2PL in these states is illustrated via sample execution. The correctness of C2V2PL scheme is proved in section 4. We conclude the paper in section 5.

2 Transaction Model

A transaction is a partial order on a set of read and write actions. The last action of the transaction, commit or abort, indicates whether its execution has completed successfully or not. Each transaction $T_i$ is assigned a unique timestamp $ts(T_i)$. For simplicity, we assume that $ts(T_i) = i$. Each action maintains the timestamp of its transaction.

We assume that the C2V2PL scheduler starts in an initial correct and consistent database state $D_0$, with a single version $x^0_0$ for each data item $x$ in the database. The notation $x^k_j$ is used as follows: $k$ is the timestamp of the transaction $T_k$ that wrote the version $x^k_j$ of the data item $x$; $j = ts(x^k_j)$ is the current timestamp of the version $x^k_j$ used in version selection to process a read action on data item $x$. As shown in the figure 1, a version for a data item $x$ is created as $x^k_k$ by the transaction $T_k$, becomes accessible to other transactions as $x^k_k$ after $T_k$ commits, and can be accessed as $x^k_0$ after $T_k$ terminates. Thus, the version $x^k_0$ of data item $x$ is always due to a terminated transaction $T_k$, and the version $x^k_j$ is always due to either active or committed but not yet terminated transaction $T_j$. We will explain the termination and commitment of a transaction later in this section.

2.1 Concurrency Control

A write action on data item $x$ in transaction $T_i$, $W_i(x)$, uses the following locking protocol.
Table 1: The higher level lock conflict matrix

<table>
<thead>
<tr>
<th></th>
<th>rl</th>
<th>w l</th>
<th>vl</th>
</tr>
</thead>
<tbody>
<tr>
<td>rl</td>
<td>✓</td>
<td>✓</td>
<td>✓</td>
</tr>
<tr>
<td>w l</td>
<td>✓</td>
<td>×</td>
<td>×</td>
</tr>
<tr>
<td>vl</td>
<td>✓</td>
<td>×</td>
<td>×</td>
</tr>
</tbody>
</table>

1. $T_i$ requests a write lock on the data item $x$.

2. scheduler grants the $wl_i(x)$ write lock on data item $x$ if there are no conflicts and the lock request satisfies the specified constraints.

3. $T_i$ creates a new version $x_i$ for the data item $x$.

As shown in Table 1, since the write locks conflict, there can be utmost one uncommitted version $x_i$ written by some transaction $T_i$, where $T_i$ holds the $wl_i(x)$ lock. As we will see later, the conflict of write lock with verified ($vl$) lock limits the number of committed versions of any data item $x$, available at a given time, to a maximum of two versions: $x_0^i$ and $x_1^i$; where the version $x_0^i$ is due to the most recently terminated transaction $T_j$ that wrote $x$ or $x_0^i$ is in initial consistent database state $D_0$, and the version $x_1^i$ is written by the currently committed but not yet terminated transaction $T_i$. The constraints that the lock request must satisfy to be granted are described in the section 3.

A read action is completed in accordance with the Read rule similar to the multi version timestamp ordering (MVTO) Read rule in [BG83].

**Read Rule:** The committed version of the data item with the largest timestamp less than or equal to the timestamp of the transaction making the read request is selected.
The scheduler maintains two versions of the read lock for each data item \( x \), i.e. \( rlo(x) \) lock and \( rl\neq 0(x) \) lock. Since utmost two committed versions of a data item are available for the scheduler to choose from, there is a one to one correspondence between the read lock version granted and the data item version selected.

A read action on a data item \( x \) in transaction \( Ti, R_i[x] \), is completed as follows.

1. \( Ti \) requests a read lock on the data item \( x \).

2. scheduler grants the \( rl^0_i(x) \) or \( rl\neq 0_i(x) \) read lock corresponding to whether the version \( x_{i} \) or version \( x^k \) (if it exists and is committed) is selected in accordance with the Read rule; and this read lock version satisfies the specified constraints.

3. \( Ti \) reads the selected version of \( x \) after obtaining the corresponding read lock version.

The scheduler processes the read action \( R_i[x] \) by selecting the committed version \( x^k \) if \( ts(T_i) \geq ts(x^k) \); and the version \( x_{i} \) otherwise; after granting the read lock version \( rl\neq 0_i(x) \) or \( rl^0_i(x) \) to \( Ti \), respectively. However, to avoid the incorrect execution as explained in Section 3, the read lock request for \( R_i[x] \) is blocked if the version \( x^k \), with \( ts(T_i) > ts(x^k) \), exists but is not committed. This lock request is said to have failed a constraint and must not be allowed to proceed. As we see in the next section, a version \( x^j \) (for some \( j \geq 0 \) always exists for each data item \( x \), which implies by the Read rule that every read action \( R_i[x] \) can be processed.

### 2.2 Version Control

We now describe the versioning control mechanism in C2V2PL scheme. A transaction can be in one of the three modes: active mode, passive mode or done mode. A transaction \( T_i \) is in active mode when it is executing its read/write actions or is blocked waiting for its
lock requests to be granted by the scheduler. A transaction $T_i$ is in passive mode when the execution of all its read and write actions has been completed successfully. A transaction $T_i$ is in done mode after the locks held by it can be released by the scheduler without compromising the future consistency of the database.

**Commit or Abort of a Transaction** The transition from the active mode to the passive mode for a transaction $T_i$ is triggered by the execution of its commit action, $c_i$. The scheduler processes the $c_i$ by converting each of the $wl_i(x)$ lock held by $T_i$ into a third kind of lock called the verified lock, $vl_i(x)$. This conversion makes the version $x^i$ written by $T_i$ accessible to the other active transactions. Thus, the commitment of a transaction represents the growing phase of the number of the committed versions of the data item written by it. None of the read locks held by $T_i$ are released during this transition. As shown in Table 1, since the $vl$ locks and the $wl$ locks conflict, no other transaction is allowed to write $x$ while $T_i$ is in passive mode, i.e. while $T_i$ is committed but has not yet terminated.

The abort action $a_i$ for transaction $T_i$ is processed by purging the new versions written by $T_i$ and releasing all the locks held by it. Since only the committed versions of any data item can be accessed by the other transactions, the cascading abort of the transactions is avoided.
**Termination of a transaction**  The transition from the passive mode to the done mode for a transaction \( T_i \) occurs when the scheduler invokes and executes the terminate action \( t_i \). The invocation of \( t_i \) determines when, for each data item \( x \) for which transaction \( T_i \) has written a committed version \( x_i^i \), can the existing \(^1\) version \( x_i^j \) be deleted so that the version \( x_i^i \) can be converted into \( x_i^j \) and the \( vl_i(x) \) lock held by \( T_i \) can be released. A transaction blocked on a write lock request on data item \( x \) can proceed only after \( T_i \) has released the \( vl_i(x) \) lock. Thus, the termination of a transaction represents the *shrinking phase* of the number of committed versions of the data items written by it. The terminate action \( t_i \) is executed as an *atomic* operation and is processed as follows.

1. the read locks held by the transaction \( T_i \) are released.

2. for each \( vl_i(x) \) lock held by \( T_i \), convert all the \( rl^\neq 0(x) \) locks, held by the other transactions in active mode, into \( rl^0(x) \) locks.

3. for each \( vl_i(x) \) lock held by \( T_i \), purge the previously existing version \( x_i^j \); convert the committed version \( x_i^j \) into a version \( x_i^0 \) by resetting the timestamp to zero; release the \( vl_i(x) \) lock.

Thus, for each data item \( x \), there exists either a terminated (and hence committed) version \( x_i^0 \) written by most recently terminated transaction \( T_j \) that wrote \( x \) or \( x_i^0 \in D_0 \); and at most one uncommitted or committed version \( x_i^j \), written by a transaction \( T_i \) that holds the exclusive \( wl_i(x) \) or \( vl_i(x) \) lock on \( x \) respectively.

It must be noted that the terminate action for transaction \( T_i \) may not be invoked immediately after \( T_i \) commits. This is because the simple assignment of the new version for every data item request in the future does not work for the reasons of consistency. Consider

\(^1\)written by most recently terminated transaction \( T_j \) that wrote \( x \) or \( x_i^0 \in D_0 \).
the two transactions \( T_1 = R_1[x] R_1[y] \) and \( T_2 = W_2[x] W_2[y] \) and the following history:

\[
D_0 \models rl_1^0(x) \cdot R_1[x_0] \cdot w_2(x) \cdot W_2[x_2^2] \cdot w_2(y) \cdot W_2[y_2^2] \cdot c_2
\]

The scheduler starts in an initial consistent database state \( D_0 \). It selects the version \( x_0^0 \) for processing \( R_1[x] \) and grants \( rl_1^0(x) \) lock to the transaction \( T_1 \). The transaction \( T_2 \) writes the versions \( x_2^2 \) and \( y_2^2 \) after it is granted the \( w_2(x) \) and \( w_2(y) \) write locks. The scheduler processes the commit action \( c_2 \) for \( T_2 \) by converting the \( w_2(x) \) and \( w_2(y) \) locks into \( vl_2(x) \) and \( vl_2(y) \) locks. The versions \( x_2^2 \) and \( y_2^2 \) become accessible to other active transactions.

Suppose that the scheduler were allowed to terminate the transaction \( T_2 \). The previously existing versions \( x_0^0 \) and \( y_0^0 \) would be deleted, and the versions \( x_2^2 \) and \( y_2^2 \) would be converted into the versions \( x_0^0 \) and \( y_0^0 \) respectively. If the scheduler now processed \( R_1[y] \) by selecting the only available version of data item \( y \), i.e. the version \( y_0^0 \) in accordance with the Read rule (since \( ts(T_1) > ts(y_0^0) \)), there would be no serial execution of the transactions \( T_1 \) and \( T_2 \). This is because in reading the version \( x_0^0 \), \( T_1 \) saw the database in a state before the execution of \( T_2 \), and in reading the version \( y_0^0 \), \( T_1 \) saw the database in a state after the execution of \( T_2 \).

To determine when the terminate action for a committed transaction can be invoked by the scheduler, we define the following irreflexive, transitive relation.

\[
T_i \text{ precedes } T_j : \iff (\exists x) [(rl_1^0(x) \text{ and } w_{lj}(x)) \text{ or } (rl_1^0(x) \text{ and } vl_{lj}(x)) \text{ or } (rl_j^{\neq0}(x) \text{ and } vl_i(x))]\]

i.e. the transaction \( T_i \) precedes the transaction \( T_j \) if either \( T_i \) has read a previously existing version of a data item for which \( T_j \) has created a new version, or \( T_j \) has read the committed version of the data item written by \( T_i \).

\[
T_j \text{ terminates : } \iff (\exists T_i) (T_i \text{ precedes } T_j)
\]

which says that the transaction \( T_j \) can not terminate, until each transaction \( T_i \) that has
either read the version $x^k_0$ (for some k) or written the committed version $x^l_1$ that has been read by $T_j$ has terminated.

By the unary relation $\text{terminates}$, in the example above, since $T_1$ has read the previously existing version $x^0_0$ of data item $x$ and $T_2$ has created the new committed version $x^2_2$, the termination of $T_2$ must be delayed until after $T_1$ has terminated. This allows the scheduler to make the correct version selection for $R_1[y]$ from the two available committed versions $y^0_g$ and $y^2_g$, i.e. the version $y^0_g$ with $ts(y^0_g) < ts(T_1) < ts(y^2_g)$. The latter requirement in $\text{terminates}$ is not as obvious and its need is illustrated with the help of another example.

Consider the following transactions $T_3 = R_3[x] R_3[y]$, $T_4 = W_4[x]$, $T_5 = R_5[x] W_5[y]$ and their execution history:

$$D_0 \mid rI^0_3(x) R_3[x^0_0] \mid wl_4(x) W_4[x^4_1] c_4 \mid rI^0_5(x) R_5[x^4_1] \mid wl_5(x) W_5[x^5_5] c_5$$

As explained above, transaction $T_4$ can not terminate until transaction $T_3$ terminates. However, suppose that transaction $T_5$ were terminated and the previously existing version $y^0_g$ replaced by version $y^5_g$ obtained from the committed version $y^5_g$. If the scheduler now processed $R_3[y]$ by selecting the only available version of data item $y$, i.e. the version $y^5_g$ in agreement with the Read rule (since $ts(T_3) > ts(y^5_g)$), there would be no serial execution of $T_3$, $T_4$ and $T_5$. $T_3$ sees the database state before $T_4$ in executing $R_3[x^0_0]$, $T_5$ sees the database state after $T_4$ in executing $R_5[x^4_1]$, and $T_3$ sees the database state after $T_5$ in executing $R_3[y^5_g]$.

It must be noted that the processing of the commit action for a transaction does not require a validation phase to check for the correctness of its execution. The execution of a transaction is guaranteed to be correct if its commit action can be submitted to the scheduler. This is because the read and write lock requests on the different versions of a data item are allowed in such a constrained manner that every read action $R_k[x^l_6]$ can
be processed in conformity with the Read rule and without leading to a non-serializable 
execution. The lock requests failing the constraints are handled in a manner concomitant 
with the scheduler state. Since the inconsistencies due to incorrect version access of a data 
item always manifest as a lock request failing the constraints, the effort in executing the 
transaction completely, only to find during the validation phase (in comparable schemes) 
that it has been executed incorrectly, can be saved by not granting such lock requests.

3 Adaptable Constrained Two Version 2PL

The C2V2PL scheme utilizes the unique timestamp associated with a transaction for or-
dering the "non conflicting" read and write lock requests on the different versions of a data 
item. It rejects or blocks the lock requests that fail to observe this ordering which is imposed 
by a set of constraints stated below. The anticipated invalidating lock requests coincide with 
these lock requests failing the constraints. It must be noted that not every such lock re-
quest will actually lead to the invalid execution of the transaction. The scheduler executes 
in one of the two states - conservative or aggressive depending on the contention for data 
among the transactions in the system. If the data contention is high, to avoid deadlocks and 
to minimize the duration for which the locks will be held by a transaction, these requests 
failing the constraints are rejected. However, if the data contention is low, to avoid the 
unnecessary abort of the transactions, these requests are blocked.

As described in the previous section, for each data item x, there is always a version $x^k_0$ 
with timestamp equal to zero, and utmost one committed version $x^j_j$ with timestamp equal 
to j. Thus, an appropriate version of data item x can always be selected for processing 
$R_i[x]$ and the corresponding read lock version can always be granted. However, a read lock 
request on a data item x by the transaction $T_i$ must satisfy the following constraint:
Constraint1: If a transaction $T_j$ holds $wl_j(x)$ lock, then $ts(T_j) \geq ts(T_i)$.

Since the transaction $T_j$ holds the $wl_j(x)$ lock, there is only one available committed \(^2\) version of data item $x$, i.e. the version $x^k_0$ for processing $R_i[x]$. Suppose this version were selected by the scheduler. If the $T_j$ commits and makes the version $x^k_0$ accessible to $T_i$, then $R_i[x]$ has not read the committed version with the largest timestamp less than $ts(T_i)$, i.e. the version $x^j_0$; hence breaking the Read rule. Thus, the lock request for $R_i[x]$ must remain blocked until it satisfies the Constraint1, i.e. until $wl_j(x)$ is converted into $vl_j(x)$ lock, or in other words until $T_j$ commits.

A write lock request $wl(x)$ for transaction $T_i$ must satisfy the following constraint:

Constraint2: There does not exist a transaction that holds $wl(x)$ or $vl(x)$ lock and for all transactions $T_j$ that hold $rl^0_j(x)$, $ts(T_i) \geq ts(T_j)$

Note that no transaction could not be holding a $rl^0(x)$ lock since no other transaction is holding a $vl(x)$ lock. This stems from the fact that terminate action always converts each of the $rl^0(x)$ locks into a $rl^0(x)$ lock before it releases its $vl(x)$ lock. The failure of Constraint2 by a write lock request may lead to the following scenario. Consider the two transactions $T_6 = R_6[x]W_6[y]$ and $T_7 = R_7[y]W_7[x]$ and the following history of execution:

$D_0 || rl^0_6(x) R_6[x_0^2] rl^0_7(y) R_7[y_0^2] wl_7(y) W_7[x_7^2] c_7$

$W_6[y]$ arrives and suppose $wl_6(y)$ lock were granted. $T_6$ now submits its commit action. The scheduler would process the request by converting the $wl_6(y)$ lock into $vl_6(x)$ lock. There is no serial execution of $T_6$ and $T_7$. But this contradicts our claim that a transaction that can submit its commit action is guaranteed to have executed correctly. The write action $W_6[x]$ is a missed write in the terminology of the MVTO scheme [BG83] and is rejected. However, in the adaptable C2V2PL scheme, such a write request may be rejected or blocked.

\(^2\)The other version $x^j_i$ written by the transaction $T_j$ is still uncommitted.
3.1 Aggressive State

The C2V2PL scheduler in the aggressive state uses the following rule for avoidance of deadlocks due to conflicting wl and vl locks:

Conflict Resolution Rule: If a transaction $T_i$ holds a $wl_i(x)$ or $vl_i(x)$ lock, then the write $wl_j(x)$ lock request $wl_j(x)$ by the transaction $T_j$ is rejected if $ts(T_j) > ts(T_i)$; and is blocked otherwise.

The conflict resolution rule along with the rejection of the write lock requests that fail the constraint makes the C2V2PL scheduler in aggressive state, deadlock free. Figure 2 shows how the timestamped lock requests are handled by the C2V2PL scheduler in the aggressive state. “X_a” and “X_b” refer to the constrained conflicting request which is rejected and blocked respectively. For example, if the transaction $T_j$ requests a $wl_j(x)$ lock when $T_i$ holds $rI^0_i(x)$, with $ts(T_i) > ts(T_j)$, the action taken by the scheduler is “X_a”, since $wl_j(x)$ lock request has failed to satisfy the Constraint. Furthermore, every read action can be completed by granting either $rI^0$ or $rI^{\neq 0}$ read lock. The action of the scheduler for the
read lock request that fails the \textit{Constraint} \textsubscript{1} is “X\textsubscript{b}”. Notice, that since the lock requests are blocked only in an asymmetric fashion, i.e. only a transaction with higher timestamp may be blocked by a lock held by a transaction with a lower timestamp, there can be no deadlocks in aggressive state of C2V2PL scheduler.

3.2 Conservative State

Based on the assumption, that in low data contention environments, there will be little inconsistent access to data, the \textit{C2V2PL} scheduler in conservative state does not reject but blocks the lock request it \textit{anticipates} will lead to an incorrect transaction execution. This lock request will result in a deadlock if its execution can \textit{indeed} lead to an invalid execution; and will be rejected when the scheduler times out to resolve this deadlock. As shown in the Fig. 3, the action of the scheduler for a write lock request that fails \textit{Constraint} \textsubscript{2} is “X\textsubscript{b}”. A transaction with a write lock request is \textit{unconditionally} blocked if another transaction already holds a write or a verified lock on that data item. The \textit{C2V2PL} scheduler in this state avoids unnecessary rejects of the lock requests that, governed by the failure of constraints are anticipated to, but do not actually lead to an incorrect execution.

Figure 3: Constrained Conflict Graph for Conservative State for C2V2PL
Illustrative Example  The following sample execution compares the behavior of the C2V2PL scheduler in aggressive and conservative state. Consider the following transactions $T_8 = R_8[z]W_8[x]$, $T_9 = R_9[z]R_9[z]W_9[y]$, $T_{10} = R_{10}[y]W_{10}[z]$ and the following order of requests submitted to the scheduler: $R_8[z], R_9[z], R_{10}[y], W_8[x], R_8[z], W_{10}[z], W_9[y]$. Assume an initial consistent database state $D_0$.

The C2V2PL scheduler in aggressive state processes $R_8[z], R_9[z],$ and $R_{10}[y]$ as $R_8[\sigma_8^0]$, $R_9[\sigma_9^0]$, and $R_{10}[\sigma_{10}^0]$ after granting the $rl_8^0(z)$, $rl_9^0(z)$, and $rl_{10}^0(y)$ locks to the transactions $T_8$, $T_9$ and $T_{10}$ respectively. The $wl_8(x)$ lock request for $W_8[x]$ fails to satisfy Constraint$_2$ and is rejected. $R_9[z]$ is processed as $R_9[\sigma_9^0]$ after the $rl_9^0(z)$ lock is granted to $T_9$. The $wl_{10}(z)$ lock request for $W_{10}[z]$ is granted and $T_{10}$ writes the version $z_{10}^0$. $T_{10}$ commits and $wl_{10}(z)$ lock is converted into $vl_{10}(z)$ lock. The $wl_9(y)$ lock request for $W_9[y]$ fails to satisfy Constraint$_2$ and is rejected. The scheduler invokes the terminate action $t_{10}$ and the $rl_{10}^0(y)$ and $vl_{10}(z)$ locks are released and the version $z_9^0$ is replaced by $z_{10}^{10}$ obtained from version $z_{10}^0$.

The C2V2PL scheduler in conservative state processes $R_8[z], R_9[z],$ and $R_{10}[y]$ in exact same way as in aggressive state. The $wl_8(x)$ lock request fails Constraint$_2$ and is blocked. $R_9[z]$ and $W_{10}[z]$ are processed as in aggressive state. $T_{10}$ commits. The $wl_9(y)$ lock request fails Constraint$_2$ and is blocked. A deadlock situation now results. To terminate $T_{10}$, the scheduler must wait until $T_9$ releases its $rl_9^0(z)$ lock. On the other hand, $T_9$ is waiting for $T_{10}$ to release its $rl_{10}(y)$, so that the $wl_9(y)$ lock request can be unblocked. The deadlock is resolved by aborting the transaction $T_9$. The $wl_8(x)$ lock request blocked by the failure of Constraint$_2$, can now be granted. $T_8$ commits and is eventually terminated by the scheduler. The scheduler can now terminate the transaction $T_{10}$.

The case of reduced number of transaction aborts in low data contention environment
at the expense of increased blocking is motivated by the C2V2PL scheduler in conservative state. In higher data contention environments, the blocking of the transactions is minimized at the cost of increased number of transaction restarts by the C2V2PL scheduler in aggressive state.

4 Correctness of C2V2PL

We will prove the correctness of C2V2PL scheme by describing it in multiversion serializability theory and confirming that all the histories produced by C2V2PL are 1SR. The interested reader is directed to the theory of multiversion serializability in [BG83].

Let H be a history over \{ T_1, T_2, T_3, \ldots \} produced by C2V2PL. Then H must satisfy the following properties.

\textbf{C2V2PL}_1: For each \( T_i \), there is a unique timestamp \( ts(T_i) \). For simplicity, we assume that \( ts(T_i) = i \).

\textbf{C2V2PL}_2: For each \( T_i \), the terminate action \( t_i \) follows the commit action, \( c_i \); i.e. \( c_i < t_i \).

\textbf{C2V2PL}_{3a}: For each \( R_k[x_0^0] \in H \), either (1) \( t_j < R_k[x_j^j] \) and \( j > 0 \); or (2) \( x_0^0 \in D_0 \).

\textbf{C2V2PL}_{3b}: For each \( R_k[x_j^j] \in H \), either (1) \( c_j < R_k[x_j^j] < t_j < t_k \) and \( ts(x_j^j) < ts(T_k) \); or (2) \( W_j[x_i^j] < R_k[x_j^j] \) and \( j = k \).

\textbf{C2V2PL}_4: For each \( R_k[x_a^i] \) and \( W_k[x_b^k] \in H \); if \( W_k[x_b^k] < R_k[x_a^i] \) then \( a = k \) and \( l = k \).

Properties \textbf{C2V2PL}_{3a,3b} together say that every Read \( R_k[x] \) either reads a committed version or reads a version written by itself (i.e. \( T_k \)). In either case, it reads the version with the timestamp less than or equal to \( ts(T_k) \). \( t_j < t_k \) in property \textbf{C2V2PL}_{3b} follows from the definition of unary relation \textit{terminates}. Property \textbf{C2V2PL}_4 says that if \( T_k \) wrote \( x \) before the scheduler received \( R_k[x] \), it translates the request to read the version written by \( T_k \).

\textbf{C2V2PL}_{5a}: For every \( R_k[x_0^0] \) and \( W_i[x_i^i] \in H \); either \( t_i < R_k[x_0^0] \) or \( R_k[x_0^0] < t_i \).
Property \(C2V2PL_{5a}\) says that \(R_k[x^0_j]\), i.e. a Read on the version \(x^0_j\), created by the terminated transaction \(T_j\), is strictly ordered with respect to the terminate action of every transaction that writes \(x\). This is because each transaction \(T_i\) that writes \(x^i_j\) holds a verified lock \(vl_i(x)\), while it waits for each transaction that has read the existing version \(x^0_j\) to terminate, before it can terminate and release \(vl_i(x)\) lock. Since the \(vl\) and \(wl\) locks conflict, for each transaction \(T_k\) that reads \(x^0_j\), either \(T_i\) must have terminated before \(T_j\) even got the \(wl_j(x)\) lock, i.e. \(t_i < wl_j(x) < t_j < R_k[x^0_j]\); or \(T_i\) must have terminated after \(T_k\) reading the version \(x^0_j\) had terminated, i.e. \(R_k[x^0_j] < t_k < t_i\).

\(C2V2PL_{5b}\): For every \(R_k[x^0_j]\) and \(W_i[x^i_j] \in H\); if \(W_i[x^i_j] < R_k[x^0_j]\) then (1) \(t_i < R_k[x^0_j]\); else (2) \(R_k[x^0_j] < t_i\) and \(t_k < t_i\).

Property \(C2V2PL_{5b}\) says that \(R_k[x^0_j]\), i.e. a Read on a committed version \(x^0_j\) due to a committed but not terminated \(T_j\) is strictly ordered with respect to the terminate action of every transaction that writes \(x\). (1) says that since the \(vl\) and \(wl\) locks conflict, \(T_i\) must have terminated and released the \(vl_i(x)\) lock before \(T_j\) even got the \(wl_j(x)\) lock, i.e. \(t_i < wl_j(x) < c_j < R_k[x^0_j]\); (2) By definition of the terminate action, \(t_j\) converts the version \(x^0_j\) read by \(R_k[x^0_j]\) into \(x^k_j\); converts the \(rl_k^0(x)\) lock into \(rl^0_k(x)\) lock; and then releases the \(wl_j(x)\) lock. By the Property \(C2V2PL_{3b}\), \(t_j < t_k\). Thus after \(T_j\) terminated and before \(T_k\) terminates, if \(ts(T_k) > ts(T_i)\), \(wl_i(x)\) lock request must wait for \(T_k\) to terminate and release the now \(rl_k^0(x)\) lock in accordance with \(Constraint_o\), i.e. \(R_k[x^0_j] < t_j < t_k < wl_i(x) < t_i\); otherwise \(T_i\) obtains the \(wl_i(x)\) lock, writes the version \(x^i_j\), and then waits for \(T_k\) that has read, the now version \(x^0_j\) to terminate. i.e. \(R_k[x^0_j] < t_j < wl_i(x) < t_k < t_i\).

\(C2V2PL_{6a}\): For every \(R_k[x^0_j]\) and \(W_i[x^i_j]\), (i, j, k distinct); if \(t_i < R_k[x^0_j]\) then \(t_i < t_j\).

\(C2V2PL_{6b}\): For every \(R_k[x^0_j]\) and \(W_i[x^0_j]\), (i, j, k distinct); if \(t_i < R_k[x^0_j]\) then \(t_i < t_j\).

Property \(C2V2PL_{6a}\) says that \(R_k[x^0_j]\) reads the most recently terminated version of \(x\).
Assume to the contrary that \( t_j < t_i \). But then, the version \( x_i^j \) generated when \( T_j \) terminated must have been deleted and replaced by \( x_0^j \) when \( T_i \) terminates, and thus \( R_k[x] \) could not have accessed \( x_i^j \). Property C2V2PL_{3b} combined with Property C2V2PL_{3b} says that \( R_k[x_j^j] \) either reads the version written by itself or the most recently committed version \( x_j^j \). Since the \( vl \) and \( wl \) locks conflict, if \( t_i < R_k[x_j^j] \) then \( t_i < W_j[x_j^j] < e_j \), which combined with Property C2V2PL_{3b} says \( t_i < t_j \).

\begin{align*}
\text{C2V2PL}_{7a}: & \quad \text{For every } R_k[x_0^i] \text{ and } W_i[x_i^i], i \neq j, j \neq k, \text{ if } R_k[x_0^i] < t_i \text{ then } t_k < t_i. \\
\text{C2V2PL}_{7b}: & \quad \text{For every } R_k[x_j^j] \text{ and } W_i[x_i^i], i \neq j, j \neq k, \text{ if } R_k[x_j^j] < t_i \text{ then } t_k < t_i. \\
\end{align*}

Property C2V2PL_{7a,7b}: says that \( T_i \) cannot terminate until every transaction that has read the existing terminated version, has terminated. Property C2V2PL_{7a} follows directly from the definition of unary relation \textit{terminates}. Property C2V2PL_{7b} follows from Property C2V2PL_{3b}.

\begin{align*}
\text{C2V2PL}_8: & \quad \text{For every } W_i[x_i^i] \text{ and } W_j[x_j^j], \text{ either } t_i < t_j \text{ or } t_j < t_i. \\
\end{align*}

Property C2V2PL_8 says that the termination of every two transactions that write the same data item are atomic with respect to each other.

**Theorem:** Every history \( H \) produced by the C2V2PL scheduler is 1SR.

**Proof:** By C2V2PL_2, C2V2PL_{3a,3b}, C2V2PL_4, \( H \) preserves reflexive reads-from relationship and is recoverable. Hence it is a MV history. Define a version order \( \ll \) as \( x^i \ll x^j \) only if \( t_i < t_j \). By C2V2PL_8, \( \ll \) is indeed a version order. We will prove that all edges in MVSG(\( H, \ll \)) are in the termination order. That is \( T_i \rightarrow T_j \) in MVSG(\( H, \ll \)) then \( t_i < t_j \).

Let \( T_i \rightarrow T_j \) be in SG(\( H \)). This edge corresponds to a reads-from relationship such as \( T_j \) reads \( x \) from \( T_i \). By C2V2PL_{3a} \( t_i < R_j[x_i^i] \) and from C2V2PL_2 \( R_j[x_0^i] < t_j \). Hence \( t_i < t_j \). Similarly, by C2V2PL_{3b} for any \( R_j[x_i^i], t_i < t_j \).
Consider a version order edge induced by $W_i[x^i], W_j[x^j]$ and $R_k[x^j_0]$, $(i, j, k$ distinct). There are two cases: $x^i \ll x^j$ or $x^j \ll x^i$. If $x^i \ll x^j$, then the version order edge is $T_i \rightarrow T_j$, and $t_i < t_j$ follows from the definition of $\ll$. If $x^j \ll x^i$, then the version order edge is $T_k \rightarrow T_i$. Since $x^j \ll x^i$, $t_j < t_i$ follows from the definition of the version order. By C2V2PL$_{6a}$ either $t_i < R_k[x^j_0]$ or $R_k[x^j_0] < t_i$. In former case, C2V2PL$_{6a}$ implies that $t_i < t_j$ contradicting $t_j < t_i$. Thus $R_k[x^j_0] < t_i$ and by C2V2PL$_{7a}$ $t_k < t_i$ as desired. The case of the version order edge induced by $W_i[x^i], W_j[x^j]$ and $R_k[x^j_0]$, $(i, j, k$ distinct) can be proved in exactly same way and is left for the reader to work out.

This proves that all edges in the MVSG(H, $\ll$) are in termination order. Since the termination order is embedded in a history, which is acyclic by definition, MVSG(H, $\ll$) is acyclic too. Thus, H is 1SR.

5 Conclusions

We have proposed a new concurrency control scheme which utilizes the versions maintained for the purpose of recovery, to allow the concurrent execution of read-write actions on different versions of a data item in a constrained manner. These constraints not only eliminates the need for validation phase in transaction execution, but in high data contention environment guarantees deadlock free execution which further reduces the lock holding times for a transaction. The constraints are specified using the unique timestamps on the transactions making the lock requests. The scheme adapts to the low data contention environments by accepting those requests that fail the constraints but do not lead to a non serializable execution.
References


