Reenactment for Read-Committed Snapshot Isolation

(Long Version)

Bahareh Sadat Arab[★], Dieter Gawlick[♦], Vasudha Krishnaswamy[♦],
Venkatesh Radhakrishnan[□], Boris Glavic[★]
[★]Illinois Institute of Technology [♦]Oracle [□]LinkedIn
{barab@hawk.,bglavic@}iit.edu, {dieter.gawlick,vasudha.krishnaswamy}@oracle.com, vradhakrishnan@linkedin.com

ABSTRACT

Provenance for transactional updates is critical for many applications such as auditing and debugging of transactions. Recently, we have introduced *MV-semirings*, an extension of the semiring provenance model that supports updates and transactions. Furthermore, we have proposed *reenactment*, a declarative form of replay with provenance capture, as an efficient and non-invasive method for computing this type of provenance. However, this approach is limited to the snapshot isolation (SI) concurrency control protocol while many real world applications apply the read committed version of snapshot isolation (RC-SI) to improve performance at the cost of consistency. We present non-trivial extensions of the model and reenactment approach to be able to compute provenance of RC-SI transactions efficiently. In addition, we develop techniques for applying reenactment across multiple RC-SI transactions. Our experiments demonstrate that our implementation in the GProM system supports efficient re-construction and querying of provenance.

1. INTRODUCTION

Tracking the derivation of data through a history of transactional updates, i.e., tracking the provenance of such operations, is critical for many applications including auditing, data integration, probabilistic databases, and post-mortem debugging of transactions. For example, by exposing data dependencies, provenance provides proof of how data was derived, by which operations, and at what time. Until recently, no solution did exist for tracking the provenance of updates run as part of concurrent transactions.

MV-semirings and Reenactment. In previous work [5], we have introduced **MV-semirings** (multi-version semirings). MV-semirings extend the semiring provenance framework [18] with support for transactional updates. We have introduced a low-overhead implementation of this model in our GProM system [6] using a novel declarative replay technique (**reenactment** [5]). This finally makes this type of provenance available to applications using the *snapshot isolation* (**SI**) concurrency control protocol. Figure 1 illustrates how reenactment is applied to retroactively compute provenance for updates and transactions based on replay with provenance capture. Consider the database states induced by a history of concurrently executed transactions. With our approach, a user can request the provenance of any transaction executed in the past, e.g., Transaction T_2 in the example. Using reenactment, a temporal query is generated that simulates the transaction's operations within the context of the transactional history and this query is instrumented for provenance capture. This so-called reenactment query is guaranteed to return the same results (updated versions of the relations modified by the transaction) as the original transaction. In the result of the reenactment query, each tuple is annotated with its complete derivation history: 1) from which previous tuple versions was it derived and 2) which updates of the transaction affected it. Importantly, reenactment only requires an audit log (a log of SQL commands executed in the past) and time travel (query access to the transaction time history of tables) to function. That is, no modifications to the underlying database system or transactional workload are required. Many DBMS including Oracle [1], DB2, and MSSQL [21] support a query-able audit log and time travel. If a system does not natively support this functionality we can implement it using extensibility mechanisms (e.g., triggers). Snapshot isolation is a widely applied protocol (e.g., supported by Oracle, PostgreSQL, MSSQL, and many others). However, the practical applicability of reenactment is limited by the fact that many real world applications use statement-level snapshots instead of transaction-level snapshots. Using statement-level snapshots improves performance and timeliness of data even though this comes at the cost of reduced consistency. In this work, we present the non-trivial extensions that are necessary to support statement-level snapshot isolation (isolation level `READ COMMITTED` in the aforementioned systems).

Snapshot Isolation (SI). Snapshot isolation [8] is a widely applied multi-versioning concurrency control protocol. Under SI each transaction T sees a private snapshot of the database containing changes of transactions that have committed before T started and T's own changes. SI disallows concurrent transactions to update the same data item. This is typically implemented by using write locks where transactions waiting for a lock have to abort if the transaction currently holding the lock commits.

Read Committed Snapshot Isolation (RC-SI). Under RC-SI each statement of a transaction sees changes of

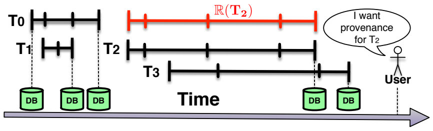


Figure 1: Reenactment

Employee

	ID	Name	Position	
$C^1_{T_0,6}(I^1_{T_0,2}(x_1))$	101	Mark Smith	Software Engineer	e_1
$C^2_{T_0,6}(I^2_{T_0,3}(x_2))$	102	Susan Sommers	Software Architect	e_2
$C^3_{T_0,6}(I^3_{T_0,4}(x_3))$	103	David Spears	Test Assurance	e_3

Bonus

	ID	EmpID	Amount	
$C^4_{T_1,10}(I^4_{T_1,8}(x_4))$	1	101	1000	b_1
$C^5_{T_2,14}(I^5_{T_2,12}(x_5))$	2	102	2000	b_2
$C^6_{T_4,18}(I^6_{T_4,16}(x_6))$	3	103	500	b_3

Figure 2: Running example database instance

Employee

	ID	Name	Position	
$C^1_{T_7,26}(U^1_{T_7,21}(C^1_{T_0,6}(I^1_{T_0,2}(x_1))))$	101	Mark Smith	Software Architecture	$e_1{}'$

Bonus

	ID	EmpID	Amount	
$C^4_{T_7,26}(U^4_{T_7,22}(C^4_{T_1,10}(I^4_{T_1,8}(x_4))))$	1	101	2000	$b_1{}'$
$C^7_{T_8,24}(I^7_{T_8,23}(C^1_{T_0,6}(I^1_{T_0,2}(x_1))))$	4	101	500	b_4

Figure 3: New and modified tuples after execution of the example history (version 26). Modified attribute values and new tuples are shown with shaded background.

	Bonus			Bonus Provenance			u_1	u_2
	ID	EmpID	Amount	P(B,ID)	P(B,EmpID)	P(B,Amount)	$\mathcal{U}_1$	$\mathcal{U}_2$
$C^4_{T_7,26}(U^4_{T_7,22}(b_1))$	1	101	2000	1	101	1000	F	T

Figure 4: Relational encoding of provenance restricted to Transaction T_7. Only tuples modified by this transaction are included and the derivation history of tuples is limited to updates of this transaction. Variables are encoded as actual tuples.

T	SQL	Time
T_7	`UPDATE Employee SET Position='Software␣Architect' WHERE ID=101;`	20
T_7	`UPDATE Bonus SET Amount = Amount + 1000 WHERE ID=101;`	21
T_8	`INSERT INTO Bonus (EmpID, Amount) (SELECT ID, 500 FROM Employee WHERE Position='Software␣Engineer');`	22
T_8	`COMMIT;`	23
T_7	`SELECT Amount INTO amounts FROM Bonus WHERE ID=101;`	24
T_7	`COMMIT;`	25

Figure 5: Example Transactional History

transactions that committed before the statement was executed. In this paper we assume the RC-SI semantic as implemented by Oracle, i.e., a statement waiting for a write-lock is restarted once the transaction holding the lock commits. This guarantees that each statement sees a consistent snapshot of the database.

Example 1. *Consider the example database shown in Figure 2 storing information about employees and the bonuses they received. Ignore the annotations to the left of each tuple for now. Two transactions have been executed concurrently (Figure 5) using the RC-SI protocol. In this example, all software engineers got a bonus of $1000 while software architects received $2000. Suppose administrator Bob executed transaction T_7 to update the position of Mark Smith to reflect his recent promotion to architect and update his bonus accordingly (increasing it by $1000). Concurrently, user Alice executed Transaction T_8 to implement the company's new policy of giving an additional bonus of $500 to all* software engineers. *All new and updated tuples for relations `Employee` and `Bonus` after the execution of these transactions are shown in Figure 3 (updated attributes are marked in red). Bob has executed a query at the end of his transaction (T_7) to double check the bonus amount for Mark expecting a single bonus of $2000 instead of the actual result (a second bonus of $500). The unexpected second bonus is produced by Transaction T_8, because this transaction did not see the uncommitted change of T_7 reflecting Mark's promotion. Thus, Mark was considered to still be a software engineer and received the corresponding $500 bonus. This kind of error is hard to debug, because it only materializes if the execution of the two transactions is interleaved in a certain way and would not occur in any serializable schedule.*

By exposing data dependencies among tuple versions (e.g., the $500 bonus for Mark is based on the previous version of Mark's tuple e_1 in the `Employee` table) and by recording which operations created a tuple version (e.g., the updated $2000 bonus for Mark was produced by the second update of T_7), the MV-semiring provenance model greatly simplifies debugging of transactions. We now give an overview of our model and then present our extensions for RC-SI.

2. THE MV-SEMIRING MODEL

Using MV-semirings (multi-version semirings), provenance is represented as annotations on tuples, i.e., each tuple is annotated with its derivation history (provenance).

$\mathcal{K}$-relations. We briefly review the semiring provenance framework [17, 18] on which MV-semirings are based on. In this framework relations are annotated with elements from an annotation domain K. Depending on the domain K, the annotations can serve different purposes. For instance, natural number annotations ($\mathbb{N}$) represent the multiplicity of tuples under bag semantics while using polynomials over a set of variables (e.g., x_1, x_2, ...) representing tuple identifiers the annotations encodes provenance. Let $\mathcal{K} = (K, +_\mathcal{K}, \times_\mathcal{K}, 0_\mathcal{K}, 1_\mathcal{K})$ be a commutative semiring. A **$\mathcal{K}$-relation** R is a (total) function that maps tuples to elements from $\mathcal{K}$ with the convention that tuples mapped to $0_\mathcal{K}$, the 0 element of the semiring, are not in the relation. A structure $\mathcal{K}$ is a commutative semiring if it fulfills the equational laws shown on the top of Figure 6. As we will see in the following, the operators of the positive relational algebra ($\mathcal{RA}^+$) over $\mathcal{K}$-relations are defined by combining input annotations using the $+_\mathcal{K}$ and $\times_\mathcal{K}$ operations where addition represents alternative use of inputs (e.g., union) and multiplication denotes conjunctive use of inputs (e.g., join). The semiring $\mathbb{N}$, the set of natural numbers with standard arithmetics corresponds to bag semantics. For example, if a tuple t occurs twice in a relation R, then this tuple would be annotated with 2 in the $\mathbb{N}$-relation corresponding to R.

Provenance polynomials. Provenance polynomials (semiring $\mathbb{N}[X]$), polynomials over a set of variables X which represent tuples in the database, model a very expressive type

Laws of commutative semirings

$$k + 0_{\mathcal{K}} = k \qquad k \times 1_{\mathcal{K}} = k \qquad \text{(neutral elements)}$$
$$k + k' = k' + k \qquad k \times k' = k' \times k \qquad \text{(commutativity)}$$
$$k + (k' + k'') = (k + k') + k'' \quad k \times (k' \times k'') = (k \times k') \times k'' \qquad \text{(associtivity)}$$
$$k \times 0_{\mathcal{K}} = 0_{\mathcal{K}} \qquad \text{(annihilation through 0)}$$
$$k \times (k' + k'') = (k \times k') + (k \times k'') \qquad \text{(distributivity)}$$

Evaluation of expressions with operands from K

$$k + k' = k +_{\mathcal{K}} k' \quad k \times k' = k \times_{\mathcal{K}} k' \quad (\text{if } k \in K \wedge k' \in K)$$

Equivalences involving version annotations

$$\mathcal{A}(0_{\mathcal{K}}) = 0_{\mathcal{K}} \qquad \mathcal{A}(k + k') = \mathcal{A}(k) + \mathcal{A}(k')$$

Figure 6: Equivalence relations for $\mathcal{K}^{\nu}$

of provenance by encoding how a query result tuple was derived by combining input tuples. Using $\mathbb{N}[X]$, every tuple in an instance is annotated with a unique variable $x \in X$ and the results of queries are annotated with polynomials over these variables. For example, if a tuple was derived by joining input tuples identified by x_1 and x_2, then it would be annotated with $x_1 \times x_2$. Since we are mainly concerned with provenance, we mostly limit the discussion to $\mathbb{N}[X]$ and its MV-semiring extension as explained below.

MV-semirings. In [5] we have introduced MV-semirings which are a specific class of semirings that encode the derivation of tuples based on a history of transactional updates. For each semiring $\mathcal{K}$, there exists a corresponding semiring $\mathcal{K}^{\nu}$, e.g., $\mathbb{N}[X]^{\nu}$ is the MV-semiring corresponding to the provenance polynomials semiring $\mathbb{N}[X]$. Since $\mathbb{N}$ encodes bag semantic relations, $\mathbb{N}^{\nu}$ represents bag semantics with embedded history. Figures 2 and 3 show examples of $\mathbb{N}[X]^{\nu}$ annotations on the left of tuples. In these symbolic $\mathbb{N}[X]^{\nu}$ expressions variables (e.g., x_1, x_2, ...) represent identifiers of freshly inserted tuples and uninterpreted function symbols called *version annotations* encode which operations (e.g., a relational update) were applied to the tuple. The nesting of version annotations records the sequence of operations that were applied to create a tuple version. For instance, consider the annotation of tuple e_1 in Figure 2. This tuple was inserted at time 2 by Transaction T_0 and was assigned an identifier 1 ($I^1_{T_0,2}$). The tuple became visible to other transactions after T_0's commit ($C^1_{T_0,6}$). Observe that these annotations encode what operations have been applied to tuples and from which other tuples they were derived.

Version Annotations. A version annotation $X^{id}_{T,\nu}(k)$ denotes that an operation of type X (one of update U, insert I, delete D, or commit C) that was executed at time $\nu - 1$ by transaction T did affected a previous version of a tuple with identifier id and previous provenance k. Assuming domains of tuple identifiers $\mathbb{I}$, version identifiers $\mathbb{V}$, and transaction identifiers $\mathbb{T}$, we use $\mathbb{A}$ to denote the set of all possible version annotations. This set contains the following version annotations for each $id \in \mathbb{I}$, $\nu \in \mathbb{V}$, and $T \in \mathbb{T}$:

$$I^{id}_{T,\nu}, U^{id}_{T,\nu}, D^{id}_{T,\nu}, C^{id}_{T,\nu} \tag{1}$$

MV-semiring Annotation Domain. In the running example, the derivation history of each tuple is a linear sequence of operations applied to a single previous tuple version. However, in the general case a tuple can depend on multiple input tuples, e.g., a query that projects an input relation onto a non-unique column ($\Pi_{Position}(Employee)$) or an update that modifies two tuples that are distinct in the input to be the same in the output (e.g., `UPDATE Employee SET ID = 101, Name = Peter`). In MV-semiring annotations this is expressed by combining the variables representing input tuples using operations $+$ and $\times$ in the expressions. Fixing a semiring $\mathcal{K}$, the domain of $\mathcal{K}^{\nu}$ is the set of finite symbolic expressions P defined by the grammar shown below where $k \in K$ and $\mathcal{A} \in \mathbb{A}$.

$$P := k \mid P + P \mid P \times P \mid \mathcal{A}(P) \tag{2}$$

For example, consider a query $\Pi_{Position}(Employee)$ evaluated over the instance from Figure 3. The result tuple `(Software Architect)` is derived from $e_1{}'$ or, alternatively, from e_2 (the two tuples with this value in attribute `position`) and, thus, would be annotated with

$$Employee(e_1{}') + Employee(e_2)$$
$$= C^1_{T_7,26}(U^1_{T_7,21}(C^1_{T_0,6}(I^1_{T_0,2}(x_1)))) + C^2_{T_0,6}(I^2_{T_0,3}(x_2))$$

We would expect certain symbolic expressions produced by the grammar above to be equivalent, e.g., expressions in the embedded semiring $\mathcal{K}$ can be evaluated using the operations of the semiring ($k_1 + k_2 = k_1 +_{\mathcal{K}} k_2$) and updating a non-existing tuple does not lead to an existing tuple ($\mathcal{A}(0_{\mathcal{K}}) = 0_{\mathcal{K}}$). This is achieved by using K^{ν}, the set of congruence classes (denoted by $[]_{\sim}$) for expressions in P based on the equivalence relations as shown in Figure 6.

Definition 1. *Let $\mathcal{K} = (K, +_{\mathcal{K}}, \times_{\mathcal{K}}, 0_{\mathcal{K}}, 1_{\mathcal{K}})$ be a commutative semiring. The MV-semiring $\mathcal{K}^{\nu}$ for $\mathcal{K}$ is the structure*

$$\mathcal{K}^{\nu} = (K^{\nu}, +_{\mathcal{K}^{\nu}}, \times_{\mathcal{K}^{\nu}}, [0_{\mathcal{K}}]_{\sim}, [1_{\mathcal{K}}]_{\sim})$$

where $\times_{\mathcal{K}^{\nu}}$ and $+_{\mathcal{K}^{\nu}}$ are defined as

$$[k]_{\sim} \times_{\mathcal{K}^{\nu}} [k']_{\sim} = [k \times k']_{\sim} \quad [k]_{\sim} +_{\mathcal{K}^{\nu}} [k']_{\sim} = [k + k']_{\sim}$$

The definition of addition and multiplication has to be read as: create a symbolic expression by connecting the inputs with $+$ or $\times$ and then output the congruence class for this expression. For example, $k = U^1_{T,\nu}(10+5)$ is a valid element of $\mathbb{N}^{\nu}$, the bag semantics MV-semiring, which denotes that a tuple with identifier 1 was produced by an update (U) of transaction T at version ν. This element k is in the same equivalence class as $U^1_{T,\nu}(15)$ based on the equivalence that enables evaluation of addition over elements from $\mathcal{K}$.

Normal Form and Admissible Instances. We have shown in [5] that $\mathcal{K}^{\nu}$ expressions admit a (non unique) normal form representing an element $k \in K^{\nu}$ as a sum $\sum_{i=0}^{n} k_i$ where none of the k_i contains any addition operations. Intuitively, each summand corresponds to a tuple under bag semantics. Thus, we will sometimes refer to a summand as a tuple version in the following. Assuming an arbitrary, but fixed, order over such summands we can address elements in such a sum by position. Following [5] we use $n(k)$ to denote the number of summands in a normalized annotation k and $k[i]$ to refer to the i^{th} element in the sum according to the assumed order. In the definition of updates we will make use of this normal form. Note that not all expressions produced by the grammar in Equation (2) can be produced by transactional histories. For instance, $U^1_{T,3}(C^1_{T,2}(\ldots)$ can never be produced by any history, because it would imply that an update of transaction T was applied after the transaction committed. An **admissible** $\mathcal{K}^{\nu}$ database instance is defined as an instance that is the result of applying a transactional history (to be defined later in this section) to an empty input database.

Example 2. *Consider the $\mathbb{N}[X]^\nu$-relation* `Bonus` *from the example shown in Figure 3. The first tuple $b_1{}'$ is annotated with $C^4_{T_7,26}(U^4_{T_7,22}\ (C^4_{T_1,10}(I^4_{T_1,8}(x_4))))$, i.e., it was created by an update of Transaction T_7, that updated a tuple inserted by T_1. Based on the outermost commit annotation we know that this tuple version is visible to transactions starting after version 25. We use the relational encoding of $\mathcal{K}^\nu$-relations from [5] restricted to tuples affected by a given transaction to be able to compute provenance using a regular DBMS and to limit provenance to a transaction of interest for a user. Figure 4 shows the relational encoding of Bonus restricted to the part of the history corresponding to transaction T_7. We abbreviate relation* `Bonus` *as B. Version annotations are represented as boolean attributes ($\mathcal{U}_i$ for update u_i) which are true if this part of the provenance has this version annotation and false otherwise. The attributes $\mathcal{U}_1$ and $\mathcal{U}_2$ represent the version annotations for the first update (u_1) and second update (u_2) of T_7. The only tuple in the instance represents the annotation of tuple (1,101, 2000). The annotation contains only a single version annotation $U^4_{T_7,22}$. Thus, only the attribute $\mathcal{U}_2$ for update u_2 corresponding to this version annotation is true and the other attribute encoding a version annotation is set to false. Variables are encoded as the input tuple annotated with the variable (b_1 in the example).*

Queries and Update Operations. We use the definition of positive relational algebra ($\mathcal{RA}^+$) over $\mathcal{K}$-relations of [5]. Let $t.A$ denote the projection of a tuple t on a list of projection expressions A and $t[R]$ to denote the projection of a tuple t on the attributes of relation R. For a condition θ and tuple t, $\theta(t)$ denotes a function that returns $1_\mathcal{K}$ if $t \models \theta$ and $0_\mathcal{K}$ otherwise.

Definition 2. *Let R and S denote $\mathcal{K}$-relations,* $\textsc{Sch}(R)$ *denote the schema of relation R, t, u denote tuples, and $k \in K$. The operators of $\mathcal{RA}^+$ on $\mathcal{K}$-relations are defined as:*

$$\Pi_A(R)(t) = \sum_{u:u.A=t} R(u) \qquad (R \cup S)(t) = R(t) + S(t)$$

$$\sigma_\theta(R)(t) = R(t) \times \theta(t) \quad \{t' \to k\}(t) = \begin{cases} k & \text{if } t = t' \\ 0_\mathcal{K} & \text{else} \end{cases}$$

$$(R \bowtie S)(t) = R(t[R]) \times S(t[S])$$
$$\text{(for any } \textsc{Sch}(R) \cup \textsc{Sch}(S) \text{ tuple } t)$$

Updates are also defined using the operations of the MV-semiring, but updates add new version annotations to previous annotations. The supported updates correspond to SQL constructs `INSERT`, `UPDATE`, and `DELETE`, and `COMMIT`. An operation is executed at a time ν as part of a transaction T. Update operations take as input a normalized, admissible $\mathcal{K}^\nu$-relation R and return the updated version of this $\mathcal{K}^\nu$-relation. An insertion $\mathcal{I}[Q, T, \nu](R)$ inserts the result of query Q into relation R. The annotations of inserted tuples are wrapped in version annotations and are assigned fresh tuple identifiers (id_{new}). An update operation $\mathcal{U}[\theta, A, T, \nu](R)$ applies the projection expressions in A to each tuple that fulfills condition θ. Both $\mathcal{U}[\theta, A, T, \nu](R)$ and $\mathcal{D}[\theta, T, \nu](R)$ wrap the annotations of all tuples fulfilling condition θ in version annotations. A commit $\mathcal{C}[T, \nu](R)$ adds commit version annotations.

Definition 3. *Let R be an admissible $\mathcal{K}^\nu$-relation. We use $\nu(u)$ to denote the version (time) when an update u was executed and $id(k)$ to denote the id of the outermost version annotation of $k \in K^\nu$. Let A be a list of projection expressions with the same arity as R, and id_{new} to denote a fresh id that is deterministically created as discussed below. Let Q be a query over a database D such that for every $\{t \to k\}$ operation in Q we have $k \in \mathcal{K}$. The update operations on $\mathcal{K}^\nu$-relations are defined as:*

$$\mathcal{U}[\theta, A, T, \nu](R)(t) = R(t) \times (\neg\theta)(t) + \sum_{u:u.A=t} \sum_{i=0}^{n(R(u))} U^{id(R(u)[i])}_{T,\nu+1}(R(u)[i]) \times \theta(u)$$

$$\mathcal{I}[Q, T, \nu](R)(t) = R(t) + I^{id_{new}}_{T,\nu+1}(Q(D)(t))$$

$$\mathcal{D}[\theta, T, \nu](R)(t) = R(t) \times (\neg\theta)(t) + \sum_{i=0}^{n(R(t))} D^{id(R(t)[i])}_{T,\nu+1}(R(t)[i]) \times \theta(t)$$

$$\mathcal{C}[T, \nu](R)(t) = \sum_{i=0}^{n(R(t))} \textsc{com}[T, \nu](k)$$

$$\textsc{com}[T, \nu](k) = \begin{cases} C^{id}_{T,\nu+1}(k) & \text{if } k = I/U/D^{id}_{T,\nu'}(k') \\ k & \text{else} \end{cases}$$

As a convention, if an attribute a is not listed in the list of expressions A of an update then $a \to a$ is assumed. For instance, abbreviating *Software Architect* as *SA* the first update of example transaction T_7 would be written as

$$\mathcal{U}[ID = 101, \text{'}SA\text{'} \to position, T_7, 20](Employee)$$

What tuple identifiers are assigned by inserts to new tuples is irrelevant as long as identifiers are deterministic and fulfill certain uniqueness requirements. Thus, we ignore identifier assignment here (see [5] for a detailed discussion).

3. CHALLENGES AND CONTRIBUTIONS

Adapting the MV-semiring model and reenactment approach to RC-SI is challenging, because the visibility rules of RC-SI are more complex than SI, i.e., different statements within a transaction can see different snapshots of the database. Under SI, the first statement of a transaction T sees a snapshot as of the time when T started and later statements see the same snapshot and the modifications of previous updates from the same transaction. Under RC-SI, each statement u also sees modifications of earlier updates from the same transaction, but in addition sees updates of concurrent transactions that committed before u executed. This greatly complicates the definition of transactional semantics in the MV-semiring model. However, as we will demonstrate it is possible to define RC-SI semantics for MV-annotated databases without extending the annotation model (only the visibility rules have to be adapted). Under SI reenactment, queries for individual updates can simply be chained together to construct the reenactment query for a transaction. However, a naive extension of this idea to RC-SI would require us to generate the version of the database seen by a certain statement u by carefully merging the snapshot of the database at the time of u's execution with the previous changes by u's transaction. Thus, while the SI reenactment query for a transaction has to read each updated relation only once, a naive approach for RC-SI would have to read each relation R once for each update that affected it. We present a solution that only has to read each relation once in most cases. Consequently, it significantly reduces the complexity of provenance computation for RC-SI transactions. The main contributions of this work are:

- We extend the **multi-version provenance model**, a provenance model for database queries, updates, and transactions to support RC-SI concurrency control protocol (Section 5).
- We extend our **reenactment** approach to support computing provenance of RC-SI workloads and present several novel optimizations that are specific to RC-SI including a technique for reducing the number of relation accesses in reenactment (Section 6).
- Our experimental evaluation demonstrates that reenactment for RC-SI is efficient and scales to large databases and complex workloads (Section 8).

4. RELATED WORK

Green et al. [17] have introduced provenance polynomials and the semiring annotation model which generalizes several other provenance models for positive relational algebra including Why-provenance, minimal Why-provenance [11], and Lineage [13]. This model has been studied intensively covering diverse topics such as relations annotated with annotations from multiple semirings [20], rewriting queries to minimize provenance [3], factorization of provenance polynomials [22], extraction of provenance polynomials from the PI-CS [15] and Provenance Games [19] models, and extensions to set difference [14] and aggregation [4]. Systems such as DBNotes [9], LogicBlox [16], Perm [15], Lipstick [2], and others encode provenance annotations as standard relations and use query rewrite techniques to propagate these annotations during query processing. Many use cases such as auditing and post-mortem transaction debugging require provenance for update operations and particularly transactions. In [5, 6], we have introduced an extension of the semiring model for SI transactional histories that is the first provenance model supporting concurrent transactions and have pioneered the reenactment approach for computing such provenance over regular relational databases. Several papers [10, 23, 7] study provenance for updates, e.g., Vansummeren et al. [23] compute provenance for SQL DML statements. This approach alters updates to eagerly compute provenance. However, developing a provenance model for transactional updates is more challenging as it requires to consider the complex interdependencies between tuple versions that are produced by concurrent transactions under different isolation levels. In this work we present the nontrivial extensions of our previous approach [5, 6] to efficiently support the RC-SI protocol which is widely used in practice.

5. READ-COMMITTED SI HISTORIES

We now define the semantics of RC-SI histories over $\mathcal{K}^{\nu}$-relations. Importantly, our extension uses standard MV-semirings and update operations. A **transaction** $T = \{u_1, \ldots, u_n, c\}$ is a sequence of update operations followed by a commit operation (c) with $\nu(u_i) < \nu(u_j)$ for $i < j$. A **history** $H = \{T_1, \ldots, T_n\}$ over a database D is a set of transactions over D with at most one operation at each version ν. We use $Start(T) = \nu(u_1)$ and $End(T) = \nu(c)$ to denote the time when transaction T did start (respective did commit). Note that the execution order of operations is encoded in the updates itself, because each update u in the MV-semiring model is associated with a version identifier $\nu(u)$ determining the order of operations.

Given a RC-SI history H we define $R[\nu]$, the annotated state of relation R at a time ν and $R[T, \nu]$, the annotated state of relation R visible to transaction T at time ν. Note that these two states may differ, because transaction T's updates only become visible to other transactions after T has committed. As in [5] we assume that histories are applied to an empty initial database. For instance, Figure 3 shows a subset of $D[26]$, the version of the example DB after execution of the history (Figure 5) over $D[18]$ (shown in Figure 2). The database state $D[18]$ is the result of running Transaction T_0 that inserted the content of the `Employee` relation and T_1, T_2, T_4 which created the tuples in relation `Bonus`.

Definition 4. *Let H be a history over a database D. The version $R[\nu]$ of relation $R \in D$ at time ν and the version $R[T, \nu]$ of relation R visible within transaction $T \in H$ at time ν are defined in Figure 7.*

Figure 7a: Relation Version in Transaction T at Time ν. To define the content of relation R at time ν within transaction T we have to distinguish between several cases: 1) per convention $R[T, \nu]$ is empty for any $\nu < Start(T)$; 2) at the start of transaction T, $R[T, \nu]$ is same as $R[\nu]$, the version of the relation containing changes of transactions committed before ν; 3) if an update was executed by transaction T at time $\nu - 1$ then its effect is reflected in $R[T, \nu]$. The update will see tuple versions created by transactions that committed before $\nu - 1$ and tuple versions created by the transaction's own updates. We use $R_{ext}[T, \nu - 1]$ to denote this version of R and explain its construction below; 4) right after transaction commit, the current version of the relation visible within T is the result of applying the commit operator to the previous version; and 5) as long as there is no commit or update on R at $\nu - 1$ then the current version of relation R is the same as the previous one.

Figure 7b: Relation Version Visible to Updates. As mentioned above we use $R_{ext}[T, \nu]$ to denote the version of relation R that is visible to an update of transaction T executed at time ν. This state of relation R contains all tuple versions created by committed transactions as long as they have not been overwritten by a previous update of transaction T (the first sum) and tuple versions created by previous updates of transaction T (the second sum). Here by overwritten we mean that a tuple version is no longer valid, because either it has been deleted or because it was updated and, thus, it has been replaced with a new updated version. Function validEx implements this check. It returns 1 if the tuple version has not been overwritten and 0 otherwise. This function uses a predicate updated(T, t, k, ν) which is true if transaction T has invalidated summand k in the annotation of tuple t before ν by either deleting or updating the corresponding tuple version. The second sum ranges over tuple versions $R[T, \nu]$ excluding tuple versions not created by transaction T (function validIn).

Figure 7c: Committed Relation Version. The committed version $R[\nu]$ of a relation R at time ν contains all changes of transactions that committed before ν. That is, all tuple versions created by any such transaction unless the

(a) Historic relation $R[T, \nu]$: version of R seen by Transaction T at Time ν

$$R[T, \nu] = \begin{cases} \emptyset & \text{if } \nu < Start(T) \\ R[\nu] & \text{if } Start(T) = \nu \\ u(R_{ext}[T, \nu - 1]) & \text{if } \exists u \in T : \nu(u) = \nu - 1 \land u \text{ updates } R \land End(T) \neq \nu - 1 \\ \mathcal{C}[T, \nu - 1](R[T, \nu - 1]) & \text{if } End(T) = \nu - 1 \\ R[T, \nu - 1] & \text{otherwise} \end{cases}$$

(b) $R_{ext}[T, \nu]$: Tuple versions visible within Transaction T at Time ν

$$R_{ext}[T, \nu](t) = \sum_{i=0}^{n(R[\nu](t))} R[\nu](t)[i] \times \textsc{validEx}(T, t, R[\nu](t)[i], \nu) + \sum_{i=0}^{n(R[T,\nu](t))} R[T, \nu](t)[i] \times \textsc{validIn}(T, t, R[T, \nu](t)[i], \nu)$$

(c) $R[\nu]$: Committed tuple versions at Time ν

$$R[\nu](t) = \sum_{T \in H \land End(T) < \nu} \sum_{i=0}^{n(R[T,\nu](t))} R[T, \nu](t)[i] \times \textsc{validAt}(T, t, R[T, \nu](t)[i], \nu)$$

(d) Validity of summands (tuple versions) within annotations

$$\textsc{validIn}(T, t, k, \nu) = 1 \text{ if } \exists \nu', k', id : k = X^{id}_{T,\nu'}(k') \land X \in \{U, D, I\}, 0 \text{ otherwise}$$

$$\textsc{validEx}(T, t, k, \nu) = 0 \text{ if } \textsc{updated}(T, t, k, \nu), 1 \text{ otherwise}$$

$$\textsc{validAt}(T, t, k, \nu) = 1 \text{ if } k = C^{id}_{T,\nu'}(k') \land (\neg \exists T' \neq T : End(T') \leq \nu \land \textsc{updated}(T', t, k, \nu)), 0 \text{ otherwise}$$

$$\textsc{updated}(T, t, k, \nu) \Leftrightarrow \exists u \in T, t', i, j : \nu(u) < \nu \land R[T, \nu(u)](t)[i] = k \land R[T, \nu(u) + 1](t')[j] = X^{id}_{T,\nu(u)+1}(k) \land X \in \{U, D\}$$

Figure 7: Historic relational instances induced by History H. $R[T, \nu]$ is the annotated instance visible by Transaction T at version ν. $R[\nu]$ is the instance containing all changes of transactions committed before version ν. Each update of a transaction sees all modifications of previous updates from the same transaction as well as modifications of transactions committed before the update was run ($R_{ext}[T, \nu]$).

tuple version is no longer valid at ν, e.g., it got deleted by another transaction. Thus, this version of relation R can be computed as the sum over all annotations on tuple t in the versions of relation R created by past transactions. However, in addition to ensuring that outdated tuple versions are not considered we also need to ensure that every tuple version is only included once. Both conditions are modelled by function $\textsc{validAt}(T, t, k, \nu)$ that return 1 if k is a summand (tuple version) in the annotation of tuple t at time ν and was created by T (this ensures that each tuple version is only added once).

Example 3. *Consider the example transactional history from Figure 5. For instance, $Bonus[T_8, 22]$ is the version of the* `Bonus` *relation seen by the insert operation of Transaction T_8 and is equal to $Bonus[22]$ (case 2, Figure 7a). It contains the tuples from the* `Bonus` *relation as shown in Figure 2, because these tuples were created by transactions that committed before time 22 (they are in $Bonus[T_8, 22]$). Thus,* $\textsc{validAt}$ *returns 1 for these tuples. For instance, tuple b_1 has been updated by Transaction T_7 (the new version is denoted as b_1') before version 22, but this transaction has not committed yet. Since T_8 has not updated b_1,* $\textsc{validEx}$ *returns 1 and the full annotation of b_1 in $Bonus[22]$ is as shown in Figure 2.*

6. REENACTMENT

We have introduced reenactment [5] as a mechanism to construct a $\mathcal{K}^\nu$-annotated relation R produced by a transaction T that is part of a history H by running a so-called reenactment query $\mathbb{R}(T)$. We have proven [5] that $\mathbb{R}(T) \equiv_{\mathbb{N}[X]^\nu} T$, i.e., the reenactment query returns the same annotated relation as the original transaction ran in the context of history H (has the same result and provenance). In this work, we present reenactment for single RC-SI transactions as well as extensions necessary to reenact a whole history. The latter requires the introduction of a operator which merges the relations produced by the reenactment queries of several transactions. This operator is also needed to compute $R_{ext}[T, \nu]$ as introduced in the previous section. After introducing this operator, we first present a method to reenact RC-SI transactions that requires merging newly committed tuples into the version of a relation visible within the reenacted transaction after every update. We then present an optimization that requires no merging in most cases and uses another new operator - version filtering.

Version Annotation Operator. For reenactment of updates and transactions we need to be able to introduce new version annotations in queries. However, the operators of $\mathcal{RA}^+$ do not support that. To address this problem, we have defined the version annotation operator in [5]. For $X \in \{I, U, D\}$ the version annotation operator $\alpha_{X,T,\nu}(R)$ takes as input a $\mathcal{K}^\nu$-relation R and wraps every summand in a tuple's annotation in $X_{T,\nu}$. The commit annotation operator $\alpha_{C,T,\nu}(R)$ only wraps summands produced by Transaction T using operator $\textsc{com}[T, \nu](k)$ from Definition 3.

$$\alpha_{X,T,\nu}(R)(t) = \begin{cases} \textsc{com}[T, \nu](k) & \text{if } X = C \\ \sum_{i=0}^{n(R(t))} X_{T,\nu}(R(t)[i]) & \text{otherwise} \end{cases}$$

Reenacting Updates. Reenactment queries for transactions are constructed from reenactment queries for single update statements. The reenactment query $\mathbb{R}(u)$ for an update u returns the modified version of the relation targeted by the update if it is evaluated over the database state seen by u's transaction at the time of the update u ($R_{ext}[T, \nu(u)]$). The semantics of update operations is the same no matter whether SI or RC-SI is applied. Thus, we can use the technique we have introduced for SI in [5] to also reenact RC-SI updates. As we will see later, it will be beneficial to let update reenactment queries operate over a different input for RC-SI than for SI which requires modifications to the update reenactment queries. Let H is a history over database D. Below we show the definitions of update reenactment queries from [5]. The reenactment query $\mathbb{R}(u)$ for operation u in H is:

$$\mathbb{R}(\mathcal{U}[\theta, A, T, \nu](R)) = \alpha_{U,T,\nu+1}(\Pi_A(\sigma_\theta(R[T, \nu]))) \cup \sigma_{\neg\theta}(R[T, \nu])$$
$$\mathbb{R}(\mathcal{I}[Q, T, \nu](R)) = R[T, \nu] \cup \alpha_{I,T,\nu+1}(Q(D[T, \nu]))$$
$$\mathbb{R}(\mathcal{D}[\theta, T, \nu](R)) = \alpha_{D,T,\nu+1}(\sigma_\theta(R[T, \nu])) \cup \sigma_{\neg\theta}(R[T, \nu])$$

For example, an update modifies a relation by applying the expressions from A to tuples that match the update condition θ. All other tuples are not affected. Thus, the result of an update can be computed as the union between these two sets. For instance, the reenactment query $\mathbb{R}(u_2)$ for the update u_2 of running example transaction T_7 is:

$$\alpha_{U,T_7,21}(\Pi_{ID,EmpID,Amount+1000 \rightarrow Amount}(\sigma_{ID=101}(Bonus[T_7, 21]))) \cup \sigma_{ID \neq 101}(Bonus[T_7, 21])$$

Transaction and History Reenactment. To reenact a transaction T, we have to connect reenactment queries for the updates of T such that the input of every update u over relation R is $R_{ext}[T, \nu(u)]$. As discussed in Section 5, this instance of relation R contains tuple versions updated by previous updates of T which targeted R as well as tuple versions from $R[\nu]$. Hence, $R_{ext}[T, \nu(u)]$ can be computed as a union between these two sets of tuples as long as we can filter out tuple versions (summands in annotations) that are no longer valid. We now introduce a new query operator that implements this filtering and then define reenactment for RC-SI transactions using this operator.

Version Merge Operator. The version merge operator $\mu(R_1, R_2)$ is used to merge two version R_1 and R_2 of a relation R such that 1) tuple versions (summands in annotations) present in both inputs are only included once in the output and 2) if both inputs include different versions of a tuple, then only the newer version is returned. This operator is used to construct $R_{ext}[T, \nu]$ from a union of $R[\nu]$ and $R[T, \nu]$. The definition of $\mu(R_1, R_2)$ is shown below.

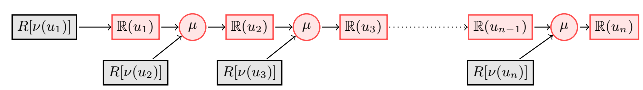
$$\mu(R_1, R_2)(t) = \sum_{i=0}^{n(R_1(t))} R_1(t)[i] \times isMax(R_2, R_1(t)[i]) + \sum_{i=0}^{n(S(t))} R_2(t)[i] \times isStrictMax(R_1, R_2(t)[i])$$

The operator uses two functions $isMax$ and $isStrictMax$. $isMax(R, k)$ returns 0 if relation R contains a newer version of the tuple version encoded as annotation k, i.e., if $\exists t', k', j : idOf(R(t')[j]) = idOf(k) \wedge versionOf(R(t')[j]) > versionOf(k)$. Function $isStrictMax$ is the strict version of $isMax$ which also returns 0 if the tuple version k is present in R, i.e., $versionOf(R(t')[j]) > versionOf(k)$ is replaced with $versionOf(R(t')[j]) \geq versionOf(k)$ in the condition. Here function $idOf(k)$ returns the tuple identifier in the annotation k and $versionOf$ returns the version encoded in the given annotation k. These functions are well defined if k is a summand in a normalized admissible $\mathcal{K}^\nu$-relation (see Section 2):

$$idOf(X^{id}_{T,\nu}(k')) = id \qquad versionOf(X^{id}_{T,\nu}(k')) = \nu$$

As an example consider computing $\mu(Bonus[26], Bonus[19])$. These relation versions are shown in Figure 2 and 3. The later only shows new or updated tuples. For instance, b_2 is present in both relations with the same annotation, a single summand. Thus, the first sum in $\mu(Bonus[26], Bonus[19])(b_2)$ will include this annotation (there is no newer version of this tuple in $Bonus[19]$) while it will be excluded from the second sum (the same annotation is found in $Bonus[26]$). As another example consider tuple b_1 which was updated to $b_1{}'$ by Transaction T_7. Thus, $\mu(Bonus[26], Bonus[19])(b_1) = 0$, because a newer version of this tuple exists in $Bonus[26]$ and $\mu(Bonus[26], Bonus[19])(b_1{}') = Bonus[26])(b_1{}')$ (this is the newest version of this tuple found in $Bonus[19]$ and $Bonus[26]$).

Reenacting Transactions. For simplicity of exposition we present the construction of reenactment queries for transactions updating a single relation R. The construction for transactions updating multiple relations is achieved analog to [5]. The reenactment query for Transaction $T = (u_1, \ldots, u_n, c)$ executed as part of an RC-SI history H is recursively constructed starting with a commit annotation operator applied to the reenactment query $\mathbb{R}(u_n)$ for the last update of T. Then we replace $R[T, \nu(u_n)]$ in the query constructed so far with $\mu(\mathbb{R}(u_{n-1}), R[\nu(u_n)])$. The result of this version merge operator is $R_{ext}[T, \nu(u_n)]$, the input seen by u_n in the history H. This replacement process is repeated for $i \in n-1, \ldots, 1$ until every reference to a version of relation R visible within the transaction has been replaced with references to committed relation versions ($R[\nu]$ for some ν). The structure of the reenactment query is outlined below.



Reducing Relation Accesses. We would like reenactment queries for RC-SI to be defined recursively without requiring to recalculate the right mix of tuple versions from transaction T and from concurrent transactions after each update. To this end we introduce the version filter operator, that filters out summands k from an annotation based on the version encoded in the outermost version annotation of k. The filter condition θ of a version filter operator is expressed using a pseudo attribute V representing the ν encoded in version annotations. We use this operator to filter summands from annotations based on the version annotations they are wrapped in.

Version Filter Operator. The version filter operator removes summands from an annotation based on the time ν in their outermost version annotation. Let θ be a condition over pseudo attribute V. Given a summand $k = X^i_{T,\nu}(k')$ such a condition is evaluated by replacing V with ν in θ. The version filter operator using such a condition θ is defined as:

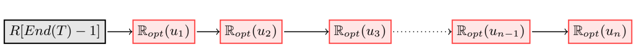
$$\gamma_\theta(R)(t) = \sum_{i=0}^{n(R(t))} R(t)[i] \times \theta(R(t)[i])$$

For example, we could use $\gamma_{V<11}(R)$ to filter out summands from annotations of tuples from a relation R that were added after time 10. In contrast to regular selection, a version filter's condition is evaluated over the individual summands in an annotation.

Our optimized reenactment approach for RC-SI is based on the following observation. Consider a tuple t updated by Transaction T and let $u \in T$ be the first update of Transaction T that modified this tuple. Let t' denote the version of tuple t valid before u. Given the RC-SI semantics, t' is obviously present in $R[\nu(u)]$ and was produced by a transaction that committed before $\nu(u)$. Importantly, t' is guaranteed to be in $R[End(T)]$, i.e, the version of R immediately before the commit of Transaction T. To see why this is the case recall that T would have obtained a write-lock on this tuple to be able to update t' to t and this write-lock is held until transaction commit. Thus, it is guaranteed that no other transaction would have been able to update t' before the commit of T. Based on this observation, we can use $R[End(T)]$ as an input to the reenactment query as long as we ensure that the reenactment queries for other updates of T executed before u ignore t'. We achieve this using the version filter operator to filter out tuple versions that were not visible to an update u'. It is applied in the input of the part of the transaction reenactment query corresponding to the update u'. In the optimized reenactment query, the initial input of reenactment is $R[End(T)]$ instead of $R[Start(T)]$. Furthermore, the update reenactment queries are modified as shown below. An optimized reenactment query $\mathbb{R}_{opt}(u)$ for update u passes on unmodified versions of tuples that are not visible to update u. We use $\mathbb{R}_{opt}(T)$ to denote the optimized transaction reenactment query. In the formulas shown below, R denotes the result of the reenactment query for the previous update or $R[End(T) - 1]$ (in case the update is the first update of the transaction). Note that this optimization is only applicable if the inserts in the transaction do not access the relation that is modified by the updates and deletes of the transaction. That is because the query of an insert may read tuple version that are not in $D[End(T)]$. Hence, we only apply this optimization if the inserts of Transaction T use the `VALUES` clause (the singleton operator $\{t \rightarrow k\}$ as defined in Section 2).

$$\begin{aligned}
\mathbb{R}_{opt}(\mathcal{U}[\theta, A, T, \nu](R)) = {} & \alpha_{U,T,\nu+1}(\Pi_A(\sigma_\theta(\gamma_{V \leq \nu(u)}(R)))) \\
& \cup\, \sigma_{\neg\theta}(\gamma_{V \leq \nu(u)}(R)) \\
& \cup\, \gamma_{V > \nu(u)}(R)) \\
\mathbb{R}_{opt}(\mathcal{D}[\theta, T, \nu](R)) = {} & \alpha_{D,T,\nu+1}(\sigma_\theta(\gamma_{V \leq \nu(u)}(R))) \\
& \cup\, \sigma_{\neg\theta}(\gamma_{V \leq \nu(u)}(R)) \\
& \cup\, \gamma_{V > \nu(u)}(R)
\end{aligned}$$

For example, the reenactment query for an update u distinguishes between three disjoint cases: 1) a tuple that is visible to the update ($V \leq \nu(u)$) and fulfills the update's condition, i.e., the tuple is updated by u; 2) a tuple that is visible to the update, but does not fulfill the condition θ; and 3) a tuple version that is not visible to u, because it was created by a transaction that committed after $\nu(u)$. The structure of the resulting reenactment query for transactions without inserts is shown below. Note that relation R is only accessed once by the reenactment query.



For each insert using the `VALUES` clause a new tuple will be added to the relation R using `UNION`.

Reenactment queries for RC-SI transactions are equivalent to the transaction they are reenacting.

Theorem 1. *Let T be a RC-SI transaction. Then, $T \equiv_{\mathbb{N}[X]^\nu} \mathbb{R}(T) \equiv_{\mathbb{N}[X]^\nu} \mathbb{R}_{opt}(T)$.*

Proof. The proof is shown in Appendix A. □

To create a reenactment query for a (partial) history, we combine the results of reenactment queries for all transactions in the history using the version merge operator. Each reference to a committed version of a relation $T[\nu]$ is replaced with a multiway merge of the results of reenactment queries for transactions $T \in H$ that committed before ν in the order of commit. For example, if two transactions T_1 and T_2 have committed before ν then $R[\nu]$ is computed as

$$q = \mu(\mathbb{R}(T_1), \mathbb{R}(T_2))$$

Later versions can then be computed by reusing this query result, e.g., if the next transaction to commit in the history was T_3, then the version of R at $End(T_3) + 1$ is computed as $\mu(q, \mathbb{R}(T_3))$.

7. IMPLEMENTATION

GProM is a middleware that implements reenactment for SI over standard DBMS using a relational encoding of MV-relations [5, 6]. Reenactment is implemented as SQL queries over this encoding. We have extended the system to implement RC-SI reenactment using the same relational encoding. One advantage of this system is that provenance requests are considered as queries and can be used as subqueries in an SQL statement, e.g., to query or store provenance. In Section 8 we study the performance of queries over provenance. GProM assumes that the underlying database system on which we want to execute provenance computations keeps an audit log that can be queried and provides at least the information as shown in Figure 5. Furthermore, the DBMS has to support time travel for the system to query past states of relations (this is used to reenact single transactions and partial histories). For instance, Oracle, DB2, and MSSQL support both features. While a full description of the implementation and additional optimizations is beyond the scope of this paper, we give a brief overview of the additional optimizations that we have implemented: 1) as we observed in [5], reenactment queries can contain a large number of union operations that may lead to bad performance if they are unfolded by the DBMS. We extend our approach for using `CASE` to avoid union operations [5] to RC-SI; 2) if the user is only interested in the provenance of tuples modified by a particular transaction, then this can be supported by filtering tuples from the output of the transaction's reenactment query that were not affected by the transaction. We did present two methods for improving the efficiency of this filter step by either removing tuples from the input of the reenactment query which do not fulfill the condition of any update of the transaction or by retrieving updated tuple versions from the database version after transaction commit and using this set to filter the input using a join. We have adapted both methods for RC-SI; 3) the version merge operator is implemented using aggregation to determine the latest version of each tuple.

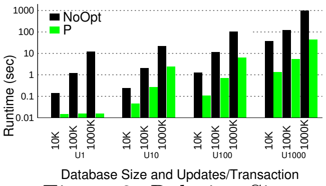


Figure 8: Relation Size

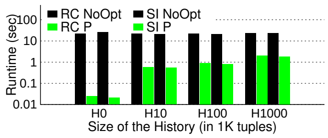


Figure 9: History Size

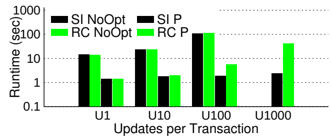


Figure 10: Isolation Levels

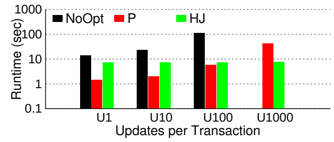


Figure 11: Optimization

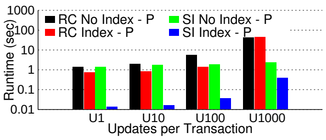


Figure 12: Index vs. No Index

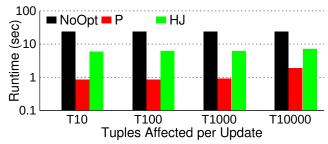


Figure 13: Affected Tuples

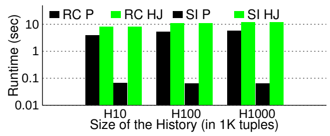


Figure 14: Inserts and Deletes

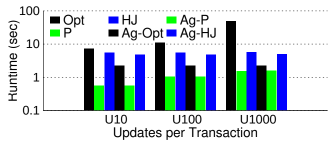


Figure 15: Aggregation

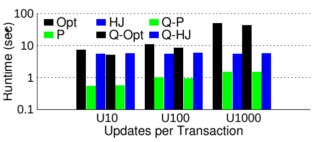


Figure 16: Query Provenance

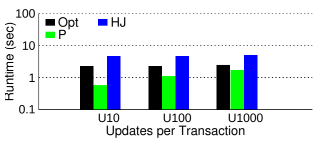


Figure 17: Query Vers. Ann.

8. EXPERIMENTS

Using commercial DBMS X, we evaluate 1) the performance of provenance computation using reenactment for isolation level *RC-SI* and comparing it with *SI*, and 2) the performance of querying provenance. All experiments were run on a machine with 2 x AMD Opteron 4238 CPUs (12 cores total), 128 GB RAM, and 4 x 1TB 7.2K HDs in a hardware RAID 5 configuration. We have studied the runtime and storage overhead of DBMS X's build-in temporal and audit features in [5]. The results demonstrated that the runtime overhead for transaction execution is below 20% when audit logging and time travel are activated and it is more efficient than eager materialization of provenance during transaction execution (about 133% overhead and higher). We did confirm the same trend for RC-SI and, thus, do not present these results here.

Datasets and Workload. In all experiments, we use a relation with five numeric columns. Values for these attributes were generated randomly using a uniform distribution. Different variants $R10K$, $R100K$, and $R1000K$ with 10K, 100K, and 1M tuples and no significant history ($H0$) were created. Moreover, three variants of $R1000K$ with different history sizes $H10$, $H100$, and $H1000$ (100K, 1M, and 10M tuples of history) are used. In most experiments, transactions consist only of update statements. The tuple to be updated is chosen randomly by its primary key. The following parameters are used in experiments: U is the number of updates per transaction (e.g., *U100* is a transaction with 100 updates). T is the number of tuples affected by each update (default is $T1$). Transactions were executed under isolation level *RC-SI* (default) or *SI*. Experiments were repeated 100 times and the average runtime is reported.

Compared Methods. We apply different configurations for computing provenance of transactions using a subset of the optimizations outlined in Section 7. **NoOpt (N)**: Computes the provenance of all tuples in a relation including tuples that were not affected by the transaction. **Opt (O)**: Like the previous option but GProM's heuristic relational algebra optimizations are activated. **Prefilter (P)**: Only returns provenance of tuples affected by the transaction by prefiltering (Section 7). **HistJoin (HJ)**: Same as *P*, but using the join method as described in Section 7.

Provenance Computation. For the following experiments we have executed the transactional workload beforehand and measure performance of provenance capture.

Relation Size and Updates/Transaction. We consider relations of different size ($R10K$, $R100K$, and $R1000K$) that do not have any significant history ($H0$). Figure 8 shows performance of computing provenance of transactions with different number of updates (*U1* up to $U1000$). We applied N and P. We scale linearly in R and U. By reducing the amount of data to be processed, the P approach is orders of magnitude faster than the N configuration.

History Size. Figure 9 shows the results for relations with 1M tuples ($R1000K$) and varying history sizes ($H0$, $H10$, $H100$, and $H1000$). We compute provenance of transactions with 10 updates (*U10*). Method N has almost constant performance for both isolation levels *RC-SI* and *SI*. The P approach displays better performance as it has to process less tuples. Its performance decreases for relations with a large history size.

Isolation Levels. Figure 10 compares the result of transactions under isolation levels *SI* and *RC-SI* with varying number of updates per transaction (*U1* to *U1000*). This experiment was conducted over table *R1000K-H1000*. The runtime of N is not affected by the choice of isolation level, because the main difference between *SI* and *RC-SI* reenactment is that we need to check whether a row version is visibile for each update. However, the impact of these checks is negligible for N as the major cost factors are scanning the table and large parts of its history as well as producing 1M output rows. For the more efficient P configuration this effect is more noticeable, especially for larger number of updates per transaction. Note that for *U1000* the N method did not finish within the allocated time budget (1000 seconds).

Comparing Optimization Techniques. Figure 11 compares different optimization methods (N, P, and *HJ*) for varying number of updates (*U1*, *U10*, *U100*, and *U1000*) using *R1000K-H1000*. Both P and *HJ* outperform N with a more pronounced effect for larger number of updates per transaction. P outperforms *HJ* for *U1* by a factor of 5 whereas this result is reversed for *U1000*. The runtime of *HJ* is almost not affected by parameter U, because it is dominated by the temporal join.

Index vs. No Index. We have studied the effect of using indexes for the relation storing the history of a relation. We use *R1000K-H1000* and vary U (*U1* to *U1000*). Figure 12 compares the effect of indexes for isolation levels *RC-SI* and *SI* using P. The results demonstrate that using indexes improves execution time of queries that apply P considerably. Provenance computation for *SI* benefits more from indexes, because the prefilter conditions applied by the P method are simpler for *SI*.

Affected Tuples Per Update. We now fix *U10* and *R1000K-H1000*, and vary the number of tuples (T) affected by each update from 10 to 10,000. The runtime (Figure 13) is dominated by scanning the history and filtering out updated tuples (P) or the self-join between historic relations (HJ). Increasing the T parameter by 3 orders of magnitude increases runtime by about 120% (P) and 9% (HJ) whereas it does not effect runtime of queries using *N*.

Inserts and Deletes. We now consider transactions that use inserts, deletes, and updates over $R1000K$ varying history size ($H10$ to $H1000$). Each statement in a transaction is chosen randomly with equal probability to be an insert, update, or delete. Figure 14 presents the result for $U20$. Performance is comparable to performance for updates for *RC-SI*. This aligns with our previous findings for *SI*.

Querying Provenance. In GProM, provenance computations can be used as subqueries of a more complex SQL query. We now measure performance of querying provenance (the runtimes include the runtime of the subquery computing provenance). All experiments of this section are run over relation $R1000K - H0$ and transactions with *U10* to *U1000*.

Aggregation of Provenance Information. Figure 15 shows the results for running an aggregation over the provenance computation (denoted as *Ag*-). These results indicate that the performance of aggregation on provenance information is comparable to provenance computation. Even more, aggregation considerably improves performance for (O). For *U1000*, *Ag-O* results in 95% improvement over O (because it reduces the size of the output) while *Ag-HJ* improves performance by $\sim 13\%$ compared to *HJ*.

Filtering Provenance. A user may only be interested in part of the provenance that fulfills certain selection conditions, e.g., bonuses larger than a certain amount. Figure 16 shows the runtime of provenance computation and querying (denoted as *Q*-). Performance of querying the results of provenance capture is actually slightly better than just computing provenance, because it reduces the size of the output and selection conditions over provenance are pushed into the SQL query implementing the provenance computation.

Querying Versions Annotations. A user can also query version annotations which are shown as boolean attributes in the provenance, e.g., to only return provenance for tuples that were updated by a certain update of the transaction. Figure 17 shows the performance results for such queries. We fix an update $u \in T$ and only return provenance of tuples modified by this update. This reduces the runtime of O queries significantly by reducing the size of the output.

9. CONCLUSIONS

We have presented an efficient solution for computing the provenance of transactions run under RC-SI by extending our MV-semiring model and reenactment approach. Our experimental evaluation demonstrates that our novel optimizations specific to RC-SI enables us to achieve performance comparable to SI reenactment. In future work, we would like to explore the application of reenactment for post-mortem debugging of transactions which is particularly important for lower isolations level such as RC-SI.

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APPENDIX

A. PROOFS

Theorem 1 *Let T be a RC-SI transaction and $\mathbb{R}(T)$ its reenactment query. Then T and $\mathbb{R}(T)$ are annotation equivalent:*

$$T \equiv_{\mathbb{N}[X]^\nu} \mathbb{R}(T)$$

Proof. Assume that transaction $T = u_1, \ldots, u_n, c$ is updating a single relation R. The proof can easily be extended for transactions updating multiple relations. We have to prove that for any potential summand k in an annotation on a tuple t, k appears in $R[T, End(T)](t)$ iff k appears in $\gamma_{\nu_e > End(xid)-1}(\mathbb{R}^R(u_n)(t))$. We prove this fact by induction over the number of updates in T.

Induction Start: Let $T = u_1, c$. We distinguish two cases. The first case applies if u_1 is an update or delete. In this case, the selection condition $V_b \leq \nu(u_1) \wedge V_e > \nu(u_1)$ in the modified reenactment query for u_1 returns $R[\nu(u_1)]$. The other input of the union is filtered using $V_b > \nu(u_1) \wedge V_e \leq \nu(u_1)$. This input may include additional summands k created by transactions which have committed between $\nu(u_1) + 1$ and $End(T)$. However, such summands will be filtered out by final version selection on condition $\gamma_{\nu_b \leq \nu(\textsc{Last}(T,R,End(T))) \wedge \nu_e > End(T)-1}$.

If u_1 is an insert, then the version selection in the modified reenactment query for u_1 ensures that the insertion query Q is evaluated over the correct input. The right input of the union simply returns $R[Start(T), End(T)-1]$ and again the final version selection filters out summands as required.

Induction Step: Assume that any transaction of length up to i is equivalent to its reenactment query. Let $T = u_1, \ldots, u_i, u_{i+1}, c$ be a RC-SI transaction of length $i+1$. We know that the induction hypothesis holds for $T_i = u_1, \ldots, u_i$, c. We have to prove that for any summand k in an annotation of a tuple t in the input of u_{i+1}, 1) iff k was updated by u_{i+1} then it will be updated by the corresponding reenactment query and 2) iff k was not updated then it will be unmodified in the result of the reenactment query. That k is present in the input of the modified reenactment query for u_{i+1} follows from the induction hypothesis. If k was updated then k would have to fulfill the version selection condition $V_b < \nu(u_i) \wedge V_e > \nu(u_i)$ and, thus, will be updated by the modified reenactment for u_{i+1}. If k was not updated then it either 1) fulfills the condition $V_b < \nu(u_i) \wedge V_e > \nu(u_i)$, but does not fulfill the condition of the update (query in case of an insert or selection condition in case of an update or delete) or 2) does not fulfill $V_b < \nu(u_i) \wedge V_e > \nu(u_i)$. In the first case, k will be in the result which follows from the correctness of update reenactment. In the second case, the modified reenactment query will ensure that k is in the result. In case of an update or delete, k will fulfill $V_b < \nu(u_i) \wedge V_e > \nu(u_i)$ in the version selection on $\neg\theta \vee V_b > \nu(u_{i+1}) \vee V_e \leq \nu(u_{i+1})$. In case of an insert the result of the modified reenactment for u_i is included unmodified (union). Since k was in the result of this query, it follows that k is in the result of $Q^R(u_{i+1})$ which concludes the proof. □

B. ADDITIONAL MV-SEMIRING BACKGROUND

B.1 Proveance Semirings

The semiring of provenance polynomials($\mathbb{N}[X]$) is the most general form of semiring annotation. The elements of this semiring are polynomials over a set of variables X which represent base tuples in the database. Usually, the assumption is that every tuple in a database instance is annotated by a unique variable $x \in X$. The provenance polynomial semiring has the important property that for any semiring $\mathcal{K}$ the annotation of a query result t in $\mathcal{K}$ can be derived from the provenance polynomial for t by mapping each variable $x \in X$ to an element from $\mathcal{K}$ and interpreting the abstract $+$ and $\times$ operations in $\mathbb{N}[X]$ as the corresponding operations in $\mathcal{K}$. For example, the semiring $\mathbb{B}$ consisting of the elements *true* and *false* using $\vee$ as addition $\wedge$ as multiplication corresponds to standard relational set semantics. The semiring $\mathbb{N}$ of the set of natural numbers with standard arithmetical operators corresponds to bag semantics. In the Lineage provenance model, the provenance of a result tuple t of a query is a set of tuples from the input were used to derive t. The semiring over the powerset of tuples in an database instance (represented as variables X) using set union for both addition and multiplication corresponds to Lineage [12].[1]

Provenance polynomials are considered the most general form of annotation in the semiring framework, because any valuation $\nu : X \to K$ of variables in X to elements from a semiring $\mathcal{K}$ can be lifted to a semiring homomorphism. Semiring homomorphisms commute with queries. This means that any type of semiring annotation can be computed from the provenance polynomial of a query result. For instance, given a query result relation with $\mathbb{N}[X]$ annotations we can compute the Lineage of the query results or bag semantics multiplicities as illustrated in the example above.

B.2 MV-semirings

The intuitive meaning of these equivalences are: 1) update operations never create tuples from non-existing or deleted tuples (recall that if a tuple is annotated with $0_{\mathcal{K}}$ in relation R this denotes that the tuple is not in the relation R) and 2) alternative use of tuples distributes over updates (e.g., updating the result of a union query returns the same result as computing the union after updating its inputs). In the following we will omit the subscript of operations and neutral elements if the semiring is clear from the context or irrelevant to the discussion. Let us state an additional equivalence that follows from the equivalences of $\mathcal{K}^\nu$: if $k' = 1_{\mathcal{K}}$ or $k' = 0_{\mathcal{K}}$, then $[\mathcal{A}(k \times k')]_\sim = [\mathcal{A}(k) \times k']_\sim$.

There exists a strong connection between $\mathcal{K}$ and $\mathcal{K}^\nu$ relations: By evaluating the version annotations (functions) in each annotation in a $\mathcal{K}^\nu$ relation we get a corresponding $\mathcal{K}$ relation. Conceptually, this means we are removing any version, transaction, and update operation information from the provenance. If we apply this approach to compute provenance polynomials from their $\mathcal{K}^\nu$ counterpart, the result will record from which tuples a tuple was derived. Later we will see that this mapping from $\mathcal{K}^\nu$ to $\mathcal{K}$ relations commutes with queries and updates, i.e., we can compute the $\mathcal{K}$ result of an operation from its $\mathcal{K}^\nu$ result (which in turns also means that this model generalizes set and bag semantics).

[1] Here $\bot$ means not in the database and $\emptyset$ means no provenance. $union_+$ and $\cup_\times$ are both standard set union except for $\bot$ where these operations are defined as $k \cup_+ \bot = \bot \cup_+ k = k$ and $k \cup_\times \bot = \bot \cup_\times k = \bot$.

Notably, the fundamental property of $\mathbb{N}[X]^\nu$, the MV-semiring of provenance polynomials, extends to update operations. The unversion operator computes not just with queries, but also with update operations.

Proposition 1. *Let $\mathcal{K}^\nu$ be an MV-semiring and V be a valuation $X \to K$. $Eval_\chi$ commute with updates.* Unv *commutes with $Eval_\chi$.*

B.3 Filtering Provenance

For databases with large histories the provenance annotation of a tuple stores its complete derivation history since the origin of the database. This amount of information can be overwhelming to a user and expensive to compute. Thus, it is important to provide a mechanism for limiting provenance information to one update operation, transaction, or a set of transactions. In the $\mathcal{K}^\nu$ model this can be achieved in a natural way by filtering parts of the annotations (to only track the effect of a certain set of statements) and by replacing subexpressions in annotations that represent parts of the history the user is not interested in with fresh variables.

Example 4. *Assume that a user is only interested the provenance of transaction T_1 from the running example. The historic database for the example history also encodes the changes applied by transaction T_2. Partial provenance for T_1 can be derived by 1) replacing subexpressions within version annotations for updates that were not executed by T_1 started with fresh variables, and 2) remove summands that were not affected by any update in T_1. In the resulting instance, only tuples affected by T_1 will be part of this partial provenance (i.e., annotated with non-zero annotations). For example, consider the annotation $C^7_{T_1,16}(D^7_{T_1,15}(C^7_{T_2,14}(I^7_{T_2,12}(x_7)))$ or m'_4 on tuple (Following, 70, 1999) in $Movie[T_1, End(T_1)]$. To limit this annotation to provenance related to T_1 we would replace $C^7_{T_2,14}(I^7_{T_2,12}(x_7))$ with a fresh variable, say y. So, its partial provenance can be indicated as $C^7_{T_1,16}(D^7_{T_1,15}(y))$.*

B.4 Reenacting SI Updates

B.5 Auxiliary Definitions For Historic Database States

In this section, we discuss the auxiliary predicates and functions which are used in the definition of historic database states based on a history H (see Figure 7). We discuss validEx(T, t, k, ν) which determines whether a summand in a normalized MV-semiring annotation that was created by a previous transaction is valid within transaction T at time ν. This function returns 1 if this part of the annotation is valid and 0 otherwise.

Analog, validIn is used

Predicate validAt(T, t, k, ν) determines whether a summand k created by a transaction T is

Determining Valid Tuple Versions. validAt(T, t, k, ν) evaluates to 1 if two conditions are met: 1) annotation k was produced by transaction T which is the case if the outermost version annotation in k is from T. Here $I/U/D$ stands for an insert, update, or delete version annotation; 2) the tuple version corresponding to k was not updated (predicate updated(T', t, k)) by another transaction T' that committed before ν ($End(T') < \nu$).

B.6 Reenacting With CASE

Proof. Consider the an input tuple t of the union in the rewritten reenactment query Q. If t fulfills $theta$ then it will appear in the left input of the union and, thus, be updated using projection A and by adding $true$ as the value of the version annotation attribute. In Q' the conditions of the conditional constructs if (θ) then e_1 else e_2 evaluate to true returning $t.A$ and adding $true$ as the value for the version annotation attribute. If t does not fulfill $theta$ then t will only appear in the right input of the union, its values will not be updated, and $false$ will be added as the value of the version annotation attribute. The same applied for Q', because the condition of the conditional constructs will evaluate to false and the original values of t's attribute values will be returned (and the version annotation attribute will be set to $false$). □

C. OPTIMIZATIONS

C.1 Prefiltering Partial Provenance

The relational encoding of reenactment queries introduced in Section 6 filters out tuples from the provenance of a transaction T that were not affected by xid by applying a selection to the result of the provenance computation that removes tuples where all version annotation are false, i.e., that are tuples that were not effected by any update of the transaction. This has the drawback the reenactment query is evaluated over all tuples from $R_{Start(T)}$. We now discuss two optimizations that filter out tuples that were not updated early on during reenactment.

C.1.1 Prefiltering using Update Conditions

The performance of the naive method can be improved if we can determine upfront which tuples will be affected by a transaction. Consider a transaction $T = (u_1, \ldots, u_n)$ and a tuple t valid at transaction start. Tuple t may be modified by a subset (potentially empty) of the updates of xid. If t is affected at all, then there has to exist a first update in T that modified tuple t. Let u_t denote this update. This first update will see the version of t that was valid at transaction start, because all update u_i with $i < t$ have not updated t. Thus, t has to fulfill the condition of u_t. This observation can be used to characterize the set of tuples affected by the transaction. In particular, the set of tuples fulfilling the condition $\theta_1 \vee \ldots \vee \theta_n$ where θ_i is the condition of the ith update operation in transaction T are exactly the tuples that where updated by T. Note that this approach is not applicable to a relation R if there exists an insert in the transaction with a query that accesses relation R. For such relations we have to fall back to the previous approach presented in Section 6.

Theorem 2 (Prefiltering). *For RC-SI transactions we can apply a technique to filter tuples from a time slice $R_{\nu_b}\nu_e$. Let T be an RC-SI transaction and R a relation affected by T that is not accessed by any insertion query in T. The rewritten reenactment query* Rew$(\mathbb{R}^R(T))$ *which applies the selection $\mathcal{A}_1 \vee \ldots \vee \mathcal{A}_n$ is equivalent to the query derived from* Rew$(\mathbb{R}^R(T))$ *by applying a selection on*

$$(\theta_1 \wedge TT_B \leq \nu(u_1) \wedge TT_e > \nu(u_1)) \vee \ldots$$
$$\vee (\theta_n \wedge TT_B \leq \nu(u_n) \wedge TT_e > \nu(u_n))$$

C.1.2 Prefiltering Using Committed Tuple Versions

Note that the version of the database at commit of transaction T will contain all tuple versions created by the transaction. Let us require that in a snapshot each tuple version has a column xid that stores which transactions created that tuple version. Thus, we can determine which tuple versions got created by a transaction T by running a query $\sigma_{xid=T}(R_{End(T)})$. To retrieve the versions of these tuples valid at transaction start (for an SI transaction), we can join the result of this query with $R_{Start(T)}$ to filter out tuple version that were updated by T. Here we assume that the database system uses some unique internal identifier for tuples that do not change between versions. This assumption holds for many DBMS. Assume that the DBMS stores these identifiers in a column *tid*. We can use this *tid* attribute to join the committed tuple versions with their counterpart at transaction start. Like the technique presented in the previous subsection this approach is only applicable to relations that are not accessed by any insertion query in the transaction.

Theorem 3 (History Join)**.** *Let T be an SI transaction and R a relation affected by T. Consider reenactment query $\mathbb{R}^R(T)$. The query* $\textsc{Rew}(\mathbb{R}^R(T))$ *is equivalent to the query derived from* $\textsc{Rew}(\mathbb{R}^R(T))$ *by replacing $R_{Start(T)}$ with*

$$\Pi_R(R_{Start(T)} \bowtie_{tid=tid'} \Pi_{tid \rightarrow tid'}(\sigma_{xid=T}(R_{End(T)})))$$

Proof. Recall that our assumption is that a *tid* attribute is not affected by any update. We know that the value of this attribute will be the same in the original version of a tuple at transaction start and at commit time. Thus, the join will return all tuples that will get updated by transaction T. □

We can adapt this technique for RC-SI transactions by using a time slice instead of a snapshot.