On the Complexity of View Update Analysis and Its Application to Annotation Propagation

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Abstract—This paper investigates three problems identified in [1] for annotation propagation, namely, the *view side-effect*, *source side-effect*, and *annotation placement problems*. Given annotations entered for a tuple or an attribute in a view, these problems ask what tuples or attributes in the source have to be annotated to produce the view annotations. As observed in [1], these problems are fundamental not only for data provenance but also for the management of view updates. For an annotation attached to a *single existing* tuple in a view, it has been shown that these problems are often intractable even for views defined in terms of simple SPJU queries [1]. We revisit these problems by considering several dichotomies: 1) views defined in various subclasses of SPJU, versus SPJU views under a practical *key preserving* condition; 2) annotations attached to existing tuples in a view versus annotations on tuples to be inserted into the view; and 3) a single-tuple annotation versus a group of annotations. We provide a complete picture of intractability and tractability for the three problems in all these settings. We show that key preserving views often simplify the propagation analysis. Indeed, some problems become tractable for certain key preserving views, as opposed to the intractability of their counterparts that are not key preserving. However, group annotations often make the analysis harder. In addition, the problems have quite diverse complexity when annotations are attached to existing tuples in a view and when they are entered for tuples to be inserted into the view.

Index Terms—Annotation, view updates, view maintenance, SPJU queries.

1 Introduction

Database annotations have been recognized by scientists as an essential feature for new generation database management systems [2], [3], [4]. Annotations are additional information attached to tuples or attributes, either entered manually or generated by programs, to explain or correct the data [2]. This information is essential to the quality and semantics of the data, and should be carried over along with the regular data when the data are migrated, transformed, or integrated. With this comes the need for studying annotation propagation. The analysis of annotation propagation is important in tracing the origin of the data [1], [2], [5], [6], [7], [8] (*a.k.a.* lineage [9], [10]), data cleaning [11], access control [12], semantic web [13], and in digital libraries [14], among other things. Several systems and tools have been developed to support annotation propagation analysis, e.g., DBNotes [15], and MONDRIAN [16].

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1.1 Annotation Propagation Analysis

In many applications, data transformations are expressed as views defined as SPJU queries in terms of the selection (S), projection (P), join (J), union (U), and renaming operators of the relational algebra. Annotations attached to some tuples in a database are carried forward to the views: the selection and projection operators preserve the annotations placed at selected tuples and the projected attributes, respectively; join merges annotations of the tuples joined together, while union simply copies the annotation of each tuple. In addition, annotating a view of some data, the annotations are carried backward to the source data as well as forward to other views [2]. As observed in [1], annotation propagation analysis is closely related to classical view update problem (see, e.g., [17] for a detailed discussion about the view update problem).

Three problems fundamental to the propagation analysis have been identified in [1], stated as follows: consider a source database D, an SPJU query Q, the view $Q(D)$, and a tuple $\triangle V$ in the view, given as input.

- The *view side-effect problem* is to find a *smallest* set $\triangle D$ of tuples in D such that $Q(D) \setminus \triangle V = Q(D \setminus \triangle D)$, i.e., with zero side effect, if such $\triangle D$ exists.
- The *source side-effect problem* is to find a smallest set $\triangle D$ of tuples in D such that $\triangle V$ is in $Q(D) \setminus Q(D \setminus \triangle D)$.
- The *annotation placement problem* is to find, given a field in the tuple $\triangle V$, a single tuple $\triangle D$ in D such that an annotation in a field of $\triangle D$ propagates to the minimum number of fields in the view including $\triangle V$.

Intuitively, when an annotation is entered for $\triangle V$ in the view, the view side-effect problem is to identify a smallest set $\triangle D$ of tuples in the source such that annotations in those places produce the view annotation, without spreading to other view tuples. Alternatively, it means that the deletion

AuName	*Journal*
Joe	TKDE
John	TKDE
Tom	TKDE
John	TODS

(a)

Journal	*Topic*	*#Papers*
TKDE	XML	30
TKDE	CUBE	30
TODS	XML	30

(b)

	AuName	*Topic*
	Joe	CUBE
	Joe	XML
	Tom	CUBE
	Tom	XML
	John	CUBE
ΔV	John	XML

(c)

AuName	*Journal*	*Topic*
Joe	TKDE	CUBE
Joe	TKDE	XML
Tom	TKDE	CUBE
Tom	TKDE	XML
John	TKDE	CUBE
John	TKDE	XML
John	TODS	XML

(d)

Fig. 1. Example of propagation problems. (a) Author(*AuName*, *Journal*). (b) Journal(*Journal*, *Topic*, *#Papers*). (c) View definition $Q_1 = \pi_{\text{AuName,Topic}}(\text{Author} \bowtie \text{Journal})$. (d) Key preserving view: $\pi_{AuName,Topic}(\text{Author} \bowtie \text{Journal})$.

of $\triangle D$ from the source leads to the removal of $\triangle V$ from the view without side effect, i.e., $\triangle D$ indicates how the tuple $\triangle V$ gets into the view.

We should remark that our statement of the view side-effect problem is referred as the "side-effect free" view side-effect problem in [1]. Here, we enforce $|\triangle V| = 0$, where $\triangle V$ denotes the difference between $Q(D)$ and $Q(D \setminus \triangle D)$. In contrast, Buneman et al. [1] aim to minimize $|\triangle V|$, although all the complexity results in [1] for the view side-effect problem are established for the "side-effect free" version, i.e., the problem studied here.

As opposed to the view side-effect problem, the source side-effect problem is to find a smallest set of tuples in the source such that the desired annotation in the view can be obtained by annotating those places in the source, although it may have side effects on the view. Note that the source side-effect problem does not require $|\triangle V| = 0$.

When some annotation is attached to a location in $\triangle V$, the annotation placement problem is to find the corresponding location in the source D to concretely annotate such that the view annotation propagates backward to the source with minimum side effects.

Example 1.1. Consider a database D with two relations: Author(*AuName, Journal*), Journal(*Journal, Topic*, *#Papers*) (with keys underlined), and an SPJ query (view definition) $Q_1 = \pi_{AuName,Topic}(\text{Author} \bowtie \text{Journal})$. Instances of both relations and the view $Q_1(D)$ are shown in Figs. 1a, 1b, and 1c (ignore Fig. 1d for now).

Suppose that John is not a researcher on XML and thus the tuple (John, XML) in the view $Q_1(D)$ is an error. Let $\triangle V = \{(\text{John}, \text{XML})\}$. We want to find tuples $\triangle D$ in the base relations of D to annotate the error such that the annotations propagate to the fields in the view tuple $\triangle V$ via Q_1; or alternatively, we want to delete $\triangle D$ such that their removal leads to the deletion of the erroneous $\triangle V$. The three problems stated above impose different conditions on how to achieve this.

1. View side-effect problem. There are multiple ways to remove tuples in D in order to delete $\triangle V$ from the view. The tuples in D related to $\triangle V$, i.e., those with matching values in $\triangle V$, are (John, TKDE), (John, TODS), (TKDE, XML, 30), and (TODS, XML, 30). We want to find a smallest set $\triangle D$ of tuples such that deleting $\triangle D$ from D leads to the removal of $\triangle V$ from $Q_1(D)$ but incurs no side effects, i.e., it deletes $\triangle V$ but no other tuples from $Q_1(D)$. To delete $\triangle V$, one can remove {(John, TKDE), (John, TODS)} from the Author table (denoted by $\triangle_1 D$), or delete (John, TKDE) from Author and (TODS, XML, 30) from Journal ($\triangle_2 D$). However, none of these is side-effect free: the first option, for example, also results in the deletion of (John, CUBE) from the view. Hence, there exists *no* solution to the view side-effect problem.
2. Source side-effect problem. It differs from condition 1 in that we do not care about the view side effect when we search for a smallest set $\triangle D$ of tuples in D to delete. Thus in this case, both $\triangle_1 D$ and $\triangle_2 D$ are solutions. In addition, removing {(TODS, XML, 30), (TKDE, XML, 30)} from Journal is also a solution although it incurs more severe view side effects than $\triangle_1 D$ and $\triangle_2 D$.
3. Annotation placement problem. Suppose that the information "John is not an XML researcher" is attached to the *AuName* field of $\triangle V$. We want to find a single tuple $\triangle D$ in the database D to annotate such that the annotation propagates to the *AuName* field of $\triangle V$ and a least number of other fields in the view $Q_1(D)$. Here, the solution is $\triangle D =$ (John, TODS): by annotating the *AuName* field of $\triangle D$, we get the desired annotation in the view with zero side effect.

1.2 View Updates and Data Provenance

The need for investigating these problems is evident in view update management and data provenance. As observed in [2], the view and source side-effect problems are the classical view deletion problems. Moreover, these two problems are important to why-provenance, while the annotation placement problem is related to where-provenance [5].

- *The connection with why-provenance.* Given an annotation attached to some tuple t in the output of a query, the view and source side-effect problems are to find which tuples in the input should be annotated such that the annotations in the input are propagated forward to the view. To answer these questions, we need to identify a (smallest) set of the input tuples that suffices to make t appear in the view. This is what why-provenance concerns, which aims to find a "proof" or "witness" for t to appear in the output, i.e., a minimum set of source tuples that suffices to produce t in the output.
- *The connection with where-provenance.* When an annotation a is attached to some field (location) l of a tuple in the view, the annotation placement problem is to find a single field (location) in the input to place the annotation a, such that a propagates to a minimum number of output locations including l. That is, we want to propagate annotations backward from the output of a query to the source database [3]. In other words, we want to identify where the value in the output location l is copied from. This is the focus of where-provenance, which is to find where a value in the output comes from.

Although there has been a host of work on annotation processing (see Section 1.4 below), the complexity bounds for the problems described above are only studied in [1], [12]. In those papers, it is shown that the analysis is in general beyond reach in practice. Indeed, the view and source side-effect problems are NP-hard for views expressed in fragments of SPJU. Similarly, the annotation placement problem is NP-hard for PJ and SPJ views [12]. While these problems are also important to the management of view updates, their complexity bounds have not been studied in that line of work.

1.3 Contributions

In this paper, we extend [1], [12] in several aspects. First, we identify a practical condition under which the analysis of annotation propagation becomes feasible. The condition, referred to as the *key preservation* condition, requires that an SPJU view Q retains a key of every base relation involved in the definition of Q. In other words, a view Q is key preserving if the primary keys of all the base relations involved in Q are included as distinct attributes in the projection fields of Q. This is less restrictive than other proposals for restricting view definitions [18], [19]. Furthermore, many views for data transformation or integration found in practice can be naturally modified to be key preserving by extending the projection-attribute list to include the primary keys.

Second, we investigate the impact of *group updates* on the analysis of annotation propagation. That is, we generalize the problem statements given earlier by allowing the given view update $\triangle V$ to include multiple tuples. The need for studying this is evident: in practice, annotations are often entered for multiple view tuples at the same time, rather than for a single tuple.

Third, in addition to annotations attached to existing tuples in a view, we study the view and source side-effect problems when the given $\triangle V$ is a set of tuples to be *inserted*. These are the classical view insertion problems. The motivation for studying this is that one often wants to know, when new tuples along with annotations are inserted into the view, how the annotations should be propagated back to the source (*a.k.a.* feedback loop [11]). We study these problems both in the presence and in the absence of the key preservation condition.

We give a full treatment of the three problems for annotation propagation *w.r.t.* the following dichotomies:

- general views versus key preserving views,
- singleton $\triangle V$ versus a set $\triangle V$ of view tuples, and
- $\triangle V$ to be deleted versus $\triangle V$ to be inserted.

We examine the impact of different combinations of these factors on the *combined complexity* of these problems. The complexity measure follows the work [1] where the complexity results of annotation propagation were first established and the studies [20], [21], [22] where the complexity of view update problems was studied.

We provide a comprehensive picture of the complexity on these problems for views defined in various fragments of SPJU queries, identifying all those cases that are intractable The results tell us the following:

1. Key preserving views often simplify the analysis of annotation propagation. For instance, the annotation placement problem is NP-hard for general PJ (with projection and join) views [12], but it is in polynomial time (PTIME) for key preserving SPJU views. When $\triangle V$ consists of a single existing tuple in the view, the source side-effect problem is NP-hard for general PJ views [1], but it is in PTIME for key preserving SPJ views. This tells us that key preserving views make it feasible to efficiently conduct certain propagation analysis.
2. Group updates complicate the propagation analysis. For instance, the view side-effect and source side-effect problems become NP-hard for key preserving SPJ views when group deletions are considered, whereas they are in PTIME for single-tuple deletions.
3. The presence of selections in the views does not complicate the analysis. More specifically, the complexity of all problems is independent of the presence of selection predicates in the view definition.
4. These problems have quite diverse complexity for view insertions and view deletions. On one hand, the view side-effect problem is in PTIME for key preserving SPU and SPJ views and single-tuple deletions, whereas it is intractable for single-tuple insertions and views defined with join only. On the other hand, the source side-effect problem is in PTIME for key preserving SPJU and single-tuple insertions, but it becomes NP-hard for JU views and single-tuple deletions.

Taken together, these tell us what cases of the annotation propagation analysis are intractable or in PTIME, for all subclasses of SPJU views, from general views to key preserving views, and from single-tuple update to group view updates. To our knowledge, no previous work has established complexity results for these problems for key preserving views, group view updates, or for view insertions. These results are useful in both the analysis of data provenance and the study of view update management.

1.4 Related Work

This paper extends an earlier version [23] as follows: 1) we investigate the annotation analysis for key preserving views that also support the union operator, and 2) we revise the statement of the view side-effect problem used in [23] to align with its counterpart in [1], and redeveloped the results accordingly. Several new intractability and PTIME results are established. In particular, we show that the view side-effect problem for insertions and views defined with join alone is already NP-hard, and that the annotation placement problem is tractable for key preserving SPJU views.

Recent research on querying annotated databases could be classified into two categories: annotation querying (e.g., [24], [25]) and annotation propagation (e.g., [1], [6], [8], [26], [27], [28]). In the former, queries access annotations as well as the regular data directly. In the latter, queries are directed primarily at the regular data, while annotations are merely carried to the query results.

More specifically, propagation schemes for processing annotations explicitly or implicitly were studies in [8], and the expressive power of various propagation schemes was investigated in [25], [26]. The problems studied there are entirely different from the problems considered in this work. Other work investigated either propagating annotations of different granularity [6], or via different methods such as semirings [27], original source tags [28], and source set [8]. The only previous complexity results on annotation propagation were established in [1], [12]. However, key preservation (to be defined in Section 2), group updates and propagation of view insertions were not considered in [1], [12].

There has also been work on modeling and managing provenance information [7], [9], [5], [29]. The focus has mostly been on different types of provenance, i.e., why-provenance [9], why-not-provenance [7], where-provenance [5], and how-provenance [29]. Only [9] gave a complexity result. In [9], a key preserving condition was also considered, which simplifies the computation of lineage. However, Cui et al. [9] studied generic mapping functions, which are quite different from SPJU views. Hence, their complexity results do not carry over to the problems considered in this paper and vice versa.

There has been a host of work on view updates. Algorithms were provided in [18] for translating restricted view updates to base table updates without side effects in the presence of certain functional dependencies. An algorithm was developed in [19] to translate (with side effects) a class of SPJ view updates to base relations, with the following restrictions: base tables may only be joined on keys and must satisfy foreign keys; a join view corresponds to a single tree where each node refers to a relation; join attributes must be preserved; and comparisons between two attributes are not allowed in selection conditions. More recently in [30], a bidirectional query language was proposed, which imposes conditions on the operators in the language such that arbitrary changes to views can be carried out. Our key preservation condition is less restrictive than those in [18], [19], [30].

An algorithm was given in [10] for translating view deletions to base relations without side effects, based on data lineage. It performs an exhaustive search over all candidate solutions, which takes exponential time. In contrast, with our key preservation condition, the computation of data lineage is simplified and the view side-effect deletion problem is PTIME resolvable. The key preservation condition was also studied in [31] for XML view updates, which is a different problem from those considered in this work.

On relational view updates, surprisingly few complexity bounds are known. The only tractability and intractability results we are aware of were established in [20], [21], [22], for finding a minimal view complement for relational views, a problem different from ours.

Commercial database systems [32], [33], [34] allow updates on very restricted views, while allowing users to specify updates manually with the INSTEAD OF triggers. For example, for views to be deletable IBM DB2 [32] restricts the from clause to reference only one base table.

1.5 Organization

We first present key preserving views in Section 2. We then establish the complexity bounds of the view side-effect problem, the source side-effect problem, and the annotation placement problem in Sections 3, 4, and 5, respectively. We identify open issues in Section 6.

Due to the space constraint, we do not include proofs that are already presented in [23]. We refer the interested reader to [23] for the details of those proofs.

2 Key Preservation

In this section, we define the notion of key preservation.

2.1 SPJU Queries

Let $\mathcal{R} = \{R_1, \ldots, R_n\}$ be a relational schema. An SPJ query (*a.k.a.* conjunctive query) on databases of $\mathcal{R}$ is a query defined in terms of the selection (σ), projection (π), join ($\bowtie$), and renaming (δ) operators in the relational algebra. It can be expressed in the normal form as follows [17]:

$$\pi_Y(R_c \times E_s), \text{ where } E_s = \sigma_F(E_c), \; E_c = S_1 \bowtie \cdots \bowtie S_n,$$

where

1. Y is a list of attributes in relations of $\mathcal{R}$;
2. $R_c = \{(A_1 : a_1, \ldots, A_m : a_m)\}$, a constant relation, such that for each $i \in [1, m]$, A_i is in Y, A_i's are distinct, and a_i is a constant in the domain of A_i;
3. for each $j \in [1, n]$, S_j is $\rho_j(R_i)$ for some R_i in $\mathcal{R}$, and ρ_j is a renaming operator, such that A_i does not appear in any S_j; and
4. F is a conjunction of equality atoms such as $A = B$ and $A =$ "a" for a constant a in the domain of A.

An SPJU query (*a.k.a.* union of conjunctive queries) defined on $\mathcal{R}$ is a query of the form $Q_1 \cup \cdots \cup Q_n$, where Q_i's are union-compatible SPJ queries on $\mathcal{R}$ [17].

We study various subclasses of SPJU, denoted by listing the operators supported. The renaming operator is included in all subclasses by default without listing it explicitly. For instance, PJ is the class of queries defined with the projection, join, and renaming operators.

For example, the view given in Fig. 1c is a PJ view.

2.2 Key Preserving Views

Consider an SPJ query $Q = \pi_Y(R_c \times E_s)$ defined above. We say that Q is *key preserving* if all the primary key attributes (with possible renaming) of each occurrence of the base relations in E_s are included in the projection fields Y of Q.

An SPJU query $Q_1 \cup \cdots \cup Q_n$ is said to be *key preserving* if for each $i \in [1, n]$, Q_i is key preserving.

Example 2.1. The query Q_1 given in Example 1.1 (and corresponding view shown in Fig. 1c) can be extended such that it is key preserving. Indeed, let $Q_2 = \pi_{AuName,Journal,Topic}(\text{Author} \bowtie \text{Journal})$. Then, Q_2 is key preserving. The view $Q_2(D)$ is shown in Fig. 1d.

The analysis of Example 1.1 becomes simpler for the key preserving view of Example 2.1. Consider the deletion of $\triangle V =$ {(John, TKDE, XML)} from $Q_2(D)$. 1) View side-effect problem. Since Q_2 is key preserving, it is obvious that the deletion can be performed by deleting either (John, TKDE) from Author or (TKDE, XML, 30) from Journal. Leveraging key preservation, we can easily check the view side effect by finding the occurrences of key values of deleted relation tuples in the view. This

TABLE 1
Complexity of the View Side-Effect Problem for Deletion Propagation

Query class	single deletion		group deletion	
	general view	key preserving view	general view	key preserving view
JU, SJU, PJU, SPJU	NP-hard ([1])	NP-hard (Cor. 2)	NP-hard ([1])	NP-hard (Cor. 2)
SPU, SP, SU, PU	PTIME ([1])	PTIME (Cor. 3)		
PJ, SPJ	NP-hard ([1])	PTIME (Prop. 1)	NP-hard ([1])	NP-hard (Thm. 4)
J, SJ	PTIME ([1])		NP-hard (Thm. 4)	

tells us that there is no solution to the problem that has zero view side effect. 2) Source side-effect problem. Similar to point 1, we can easily determine that the solution is either {(John, TKDE)} or {(TKDE, XML, 30)}. 3) Annotation placement problem. Similarly, we can see that the solution is {(John, TKDE)}.

3 The View Side-Effect Problem

In this section, we investigate the view side-effect problem. We first study the problem for single-tuple and *group deletions* in Section 3.1. We then investigate the problem for *insertions* in Section 3.2, for key preserving SPJU views and general SPJU views.

3.1 Deletion Propagation

Given a view deletion $\triangle V$, a source database D, an SPJU query Q, and the view $Q(D)$, the view side-effect problem for deletion propagation is to find a smallest set of source tuples $\triangle D$ to delete such that tuples in $\triangle V$ are deleted, without side effects.

The view side-effect problem has been studied in [1] for general SPJU views and single-tuple deletions (when $\triangle V$ is a singleton set). We investigate this problem for key preserving SPJU views and single-tuple deletions, and for SPJU views (key preserving or not) and group deletions (when $\triangle V$ consists of multiple tuples).

We present in Table 1 the complexity of the view side-effect problem for various subclasses of SPJU, for single-tuple deletions, and for group deletions (see Section 6).

3.1.1 Single-Tuple Deletions

It is known that without the key preservation condition, the view side-effect problem for single deletions on a PJ view is NP-hard [1]. In contrast, the problem becomes tractable for key preserving SPJ views. This shows that key preservation may indeed simplify the analysis of annotation propagation.

Proposition 1. *The view side-effect problem is in PTIME for key preserving SPJ views and single-tuple deletions.*

Proof Sketch. We show that for key preserving SPJ views, it suffices to delete a view tuple by removing any component tuple from a base relation that "contributes to" the view tuple. Moreover, it is in PTIME to check if the deletion is side-effect free. We refer the interested reader to [23, Theorem 3.1] for a detailed proof. □

However, key preservation does not make our lives easier for JU views. From the proof in [1] for JU views, it follows that the problem remains NP-hard for key preserving JU views.

Corollary 2. *The view side-effect problem is NP-hard for key preserving JU views and single-tuple deletions.*

Proof. The proof of [1, Theorem 2.2] is applicable here. The proof shows that the view side-effect problem is NP-hard for single-tuple deletions and JU views by reduction from the 3SAT problem. The reduction, however, uses JU views that are key preserving. □

3.1.2 Group Deletions

Our first result for group deletions is a PTIME algorithm for the view side-effect problem for SPU views. It should be remarked that the complexity of group view deletions is considered in neither [1] nor [12].

Corollary 3. *The view side-effect problem is in PTIME for* SPU *views and for group deletions.*

Proof. Let $\mathcal{R}$ be a schema and D database as described in Proposition 1, $Q = \bigcup_{j=1}^{k} Q_j$ a union of SP queries, and $\triangle V = \{t_1, \ldots, t_m\}$ be a group deletion. We give a PTIME algorithm for computing $\triangle D$ that is side-effect free, if it exists. The algorithm is an extension of the algorithm for single deletions developed in [1].

The algorithm first scans D and returns for each $R_i \in \mathcal{R}$ the set of tuples satisfying at least one of the selection conditions in one of the SP queries Q_j. We denote the resulting database by D'_0. Next, the algorithm considers the tuples in $\triangle V$ and removes them from $\triangle V$ if a side-effect free update has been found for the tuples considered so far. That is, at each step, a set $\triangle D'$ is computed, if it exists, such that $Q(D \setminus \triangle D') = Q(D) \setminus \triangle V'$, where $\triangle V'$ denotes the tuples in $\triangle V$ removed so far. It is important to remark that for SPU views, the set $\triangle D'$, if it exists, is uniquely determined.

Initially, $D' = D'_0$, $\triangle V' = \emptyset$, and $\triangle D = \emptyset$. As long as $\triangle V \setminus \triangle V' \neq \emptyset$, let t be the first tuple (based on some arbitrary ordering) in $\triangle V \setminus \triangle V'$. The algorithm then computes $\triangle D'$ as the set of all tuples from D' that project on t. Observe that this set is indeed uniquely determined. We distinguish between the following two cases: 1) there exists a tuple $s \in \triangle D'$ such that $Q(\{s\}) \not\subseteq \triangle V$. In this case, the algorithm halts and no side-effect free solution exists. 2) $Q(\triangle D') \subseteq \triangle V$. In this case, the deletion of $\triangle D'$ in D' causes side effects that all belong to $\triangle V$. Clearly, these view tuples do not have to be further processed by the algorithm. Hence, we set $\triangle D = \triangle D \cup \triangle D'$ and call the algorithm with $D' = D' \setminus \triangle D'$ and $\triangle V' = \triangle V' \cup (\triangle V \cap Q(\triangle D'))$. If the algorithm successfully terminates, i.e., when $\triangle V' = \triangle V$, the side-effect free update $\triangle D$ is returned. The solution $\triangle D$ is minimum because of the uniqueness of each update computed during the execution of the algorithm, and in

TABLE 2
Complexity of the View Side-Effect Problem for Insertion Propagation

Query class	single insertion		group insertion	
	general view	key preserving view	general view	key preserving view
PU, SPU	NP-hard (Thm. 6)			
J, SJ, PJ, JU, SPJ, SJU, PJU, SPJU	coNP-hard (Prop. 5)			
PU,SU	PTIME (Thm. 7)			

addition, the need to remove all tuples in these updates that is necessary in any solution. The algorithm clearly runs in polynomial time. □

Group updates may complicate the annotation propagation analysis. In contrast to Proposition 1, the view side-effect problem becomes NP-hard for group deletions and key preserving views defined with join only.

Theorem 4. *The view side-effect problem is NP-hard for key preserving J views and group deletions.*

Proof Sketch. We prove this by reduction from the minimum set cover problem. The detail can be found in the proof of [23, Theorem 3.3]. □

3.2 Insertion Propagation

Given a source database D, a query Q, the view $V = Q(D)$, and a set $\triangle V$ of tuples, the view side-effect problem for insertion propagation is to find a minimum set $\triangle D$ of tuples such that $Q(D \cup \triangle D) = Q(D) \cup \triangle V$, i.e., the insertion of $\triangle D$ into D produces $\triangle V$ and does not incur any side effect.

For single-tuple and group insertions, the complexity bounds of the view side-effect problem are summarized in Table 2 for various fragments of SPJU views, key preserving or not. Compared with Table 1, one can see that view insertions complicate the view side-effect analysis. In particular, the problem becomes coNP-hard for key preserving views defined with join only, even for single-tuple insertions.

To see the complication introduced by view insertions, consider the differences between insertions and deletions for key preserving views. To insert a tuple t into V, one can identify the key k_i of the tuple t_i that needs to be inserted into each occurrence of each R_i relation involved in V. As will be seen shortly, based on k_i, one can either identify an existing tuple t_i in the R_i relation with k_i, or otherwise, construct a tuple t_i carrying k_i as its key and insert it into the R_i relation. Observe that while view tuple deletions can always be carried out when side effects are allowed, it is not always doable to insert a tuple into views in the presence of the key preservation condition, even if side effects are allowed.

Example 3.1. Consider the key preserving query $Q_3 = (\text{Author} \bowtie \text{Journal})$ in the setting of Example 1.1, and the insertion of tuple (Kate, TODS, XML, 35) into the view $Q_3(D)$. At first glance, it seems that this insertion can be carried out by inserting (Kate, TODS) into table Author and (TODS, XML, 35) into Journal. However, this insertion is not possible: the insertion of (TODS, XML, 35) has to be rejected since taken together with (TODS, XML, 30) it violates the key in the table Journal.

3.2.1 Intractability Results

In contrast to Proposition 1, the view side-effect problem is already intractable when Q is a key preserving view defined with join only, even if $\triangle V$ consists of a single tuple to be inserted.

Proposition 5. *The view side-effect problem is coNP-hard for J views and single-tuple insertions.*

Proof. We prove the coNP-hardness by reduction from the complement of the Boolean conjunctive query problem. An instance of that problem is an SJ query $q = \sigma_C(\delta_{f_1}(R_1) \bowtie \cdots \bowtie \delta_{f_k}(R_k))$ over an instance I of relational schema $\mathcal{R} = \{R_1, \ldots, R_m\}$. It is to decide whether $q(I) \neq \emptyset$. This problem is NP-complete (cf. [35]).

Given q and I, we define a source database D, a J query Q, and a single tuple $\triangle V$ to be inserted into the view $V = Q(D)$, such that $q(I) = \emptyset$ iff there exists a side-effect free solution $\triangle D$.

Base relations. Database D consists of $1 + m + w$ relations, where w is the number of the selection conditions in C of the form "$\sigma_{A=c}$," for some attribute A and constant c. More specifically, the database includes 1) $R_0(A)$, initially empty, where A is a distinct attribute; 2) R'_1, and R_i for $i \in [2, m]$ to code the input instance I, where R'_1 is R_1 extended with the attribute A; and 3) a new relation S_i with one attribute A_i for each selection condition "$\sigma_{A_i=c}$" in C, to encode constants in the selection condition C. Initially, $S_i = \{(c)\}$.

View. We first define a J query $Q' = R_0 \bowtie q' \bowtie S_1 \bowtie \cdots \bowtie S_w$, where q is obtained from q by replacing each occurrence of R_1 by R'_1. Suppose that C contains p selection conditions "$\sigma_{A_i=B_i}$" for some attributes A_i and B_i, for $i \in [1, p]$. We incorporate these equality conditions into the view one by one. Initially, $Q_0 = Q'$. Suppose that we have already encoded the first $j-1$ conditions as Q_{j-1}. Let $\sigma_{A_j=B_j}$ be the next selection condition. We then define $Q_j = Q_{j-1} \bowtie \delta_{A_j/B_j}(q')$, where δ_{A_j/B_j} renames A_j as B_j, while keeping the other attributes unchanged. Finally, we define $Q = Q_p$. Note that the size of Q is quadratic in the size of q. Initially, $Q(D) = \emptyset$.

View insertions. We define $\triangle V$ as a single tuple $(x, x, \ldots, x)$ to be inserted into $Q(D)$, where x is a distinct value.

See Fig. 2 for an illustration of the reduction.

One can readily verify that $q(I) = \emptyset$ iff there exists a (minimum) $\triangle D$ such that $Q(D \cup \triangle D) = V \cup \triangle V$. □

As opposed to Corollary 3, the problem also becomes harder for key preserving PU views and insertions. The intractability remains intact on general PU views.

Theorem 6. *The view side-effect problem is NP-hard for key preserving* PU *views and single-tuple insertions.*

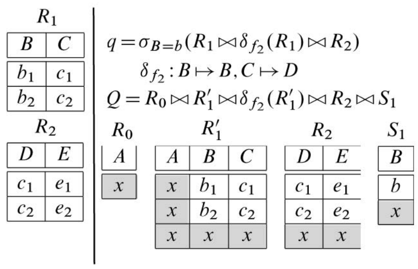

R_1
$q = \sigma_{B=b}(R_1 \bowtie \delta_{f_2}(R_1) \bowtie R_2)$
$\delta_{f_2} : B \mapsto B, C \mapsto D$
$Q = R_0 \bowtie R'_1 \bowtie \delta_{f_2}(R'_1) \bowtie R_2 \bowtie S_1$
R_2
R_0
R'_1
S_1

Fig. 2. Illustration of the proof of Proposition 5.

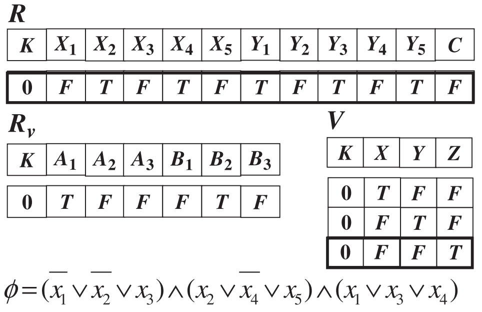

R
R_v
V
$\phi = (\bar{x}_1 \vee \bar{x}_2 \vee x_3) \wedge (x_2 \vee \bar{x}_4 \vee x_5) \wedge (x_1 \vee x_3 \vee x_4)$

Fig. 3. Illustration of the proof of Theorem 6.

Proof. We prove the NP-hardness by reduction from the 1-in-3 3SAT problem. An instance of the latter is $\phi = C_1 \wedge \cdots \wedge C_n$, where all variables in ϕ are $x_1, \ldots, x_k$, and each clause C_j is of the form $\ell_{j_1} \vee \ell_{j_2} \vee \ell_{j_3}$ and ℓ_{i_j} is either x_s or $\bar{x}_s$, $s \in [1, k]$. The problem is to determine whether there is a truth assignment that makes ϕ true and for which exactly one literal in each clause is assigned true. This problem is NP-complete (cf. [35]).

Given ϕ, we define a source database D, a key preserving PU query Q, and a single tuple $\triangle V$ to be inserted into the view $V = Q(D)$, such that ϕ has a 1-in-3 truth assignment iff there exists a minimum $\triangle D$ that is side-effect free, i.e., $Q(D \cup \triangle D) = V \cup \triangle V$.

Base relations. The database D consists of two relations $R(K, X_1, \ldots, X_k, Y_1, \ldots, Y_k, C)$ and $R_v(K, A_1, A_2, A_3, B_1, B_2, B_3)$. Here, K is the key attribute of R and is to enforce that zero-side-effect solutions consist of a single-tuple insertion in R only. The attributes X_i and Y_i are to encode truth values $\{T, F\}$, and C is an auxiliary attribute needed to check that tuples in R are truth assignments of $X = \{x_1, \ldots, x_k\}$. Initially, R is empty. As will be seen shortly, the view will extract all permutations of the truth values of the literals occurring in each clause. Clearly, when dealing with 1-in-3 truth assignments, this set of permutations is necessarily limited to $\{(T, F, F), (F, T, F), (F, F, T)\}$. However, since our view is initially empty and we are considering single-tuple insertions only, we use the relation R_v to populate the initial view with two fixed tuples, corresponding to the permutations $\{(T, F, F), (F, T, F)\}$. For this, R_v consists of a single tuple $(0, T, F, F, F, T, F)$. As before, the key attribute K of R_v is to avoid the insertion of additional tuples in R_v.

View. We define a key preserving PU query $Q = V_0 \cup V_1 \cup V_2$ as follows:

- $V_0 = \pi_{K,X,Y,Z}(\delta_{f_1}(R_v)) \cup \pi_{K,X,Y,Z}(\delta_{f_2}(R_v))$, where δ_{f_1} renames A_1, A_2, A_3 as X, Y, Z, and δ_{f_2} renames B_1, B_2, B_3 as X, Y, Z. This query yields two tuples corresponding to two of the three valid 1-in-3 truth assignments of clauses. Initially, $V_0(D) = \{(0, T, F, F), (0, F, T, F)\}$.
- $V_1 = \bigcup_{i=1}^{k} \bigcup_{\delta \in \lambda(X_i, Y_i, C)} \pi_{K,X,Y,Z}(\delta(R))$, where $\lambda(X_i, Y_i, C)$ is the set of all (bijective) renaming of X_i, Y_i, C into X, Y, Z. Intuitively, each such renaming corresponds to a permutation of X_i, Y_i, C. Since C will be enforced to equal to F (false), the view V_1 is used to verify whether tuples in R define a truth assignment. Indeed, μ defines a truth assignment if the permutations of $(\mu(x), \mu(\bar{x}), F)$ correspond to $\{(T, F, F), (F, T, F), (F, F, T)\}$, i.e., the tuples in the view. Initially, $V_1(D) = \emptyset$.
- $V_2 = \bigcup_{i=1}^{n} V_{2,i}$, where each $V_{2,i}$ is to encode the clause C_i of ϕ. Suppose that $C_i = \ell_{j_1} \vee \ell_{j_2} \vee \ell_{j_3}$. If $\ell_{i_j} = x_s$ for some $s \in [1, k]$, then let $A_{i_j} = X_s$; if $\ell_{i_j} = \bar{x}_s$, then $A_{i_j} = Y_s$. We then define $V_{2,i} = \bigcup_{\delta \in \lambda(A_{j_1}, A_{j_2}, A_{j_3})} \pi_{K,X,Y,Z}(\delta(R))$, where as before, $\lambda(A_{j_1}, A_{j_2}, A_{j_3})$ is the set of all (bijective) renaming of $A_{j_1}, A_{j_2}, A_{j_3}$ into X, Y, Z. Initially, $V_2(D) = \emptyset$.

This view simply checks whether all permutations of literals in all clauses conform the 1-in-3 condition. Note that Q is a key preserving PU query.

View insertions. We define $\triangle V$ to consist of a single tuple $(0, F, F, T)$ to be inserted into $V = Q(D)$.

The reduction is illustrated in Fig. 3 for $\phi = (\bar{x}_1 \vee \bar{x}_2 \vee x_3) \wedge (x_2 \vee \bar{x}_4 \vee x_5) \wedge (x_1 \vee x_3 \vee x_4)$. The tuple inserted into the view V and the tuple to be inserted into D are indicated by the bold rectangles.

We next verify that there is a 1-in-3 truth assignment for ϕ iff there exists a minimum side-effect free $\triangle D$.

First, assume that μ is a 1-in-3 truth assignment for ϕ. Let $\triangle D$ be the single tuple $(0, \mu(x_1), \mu(x_2), \ldots, \mu(x_n), \mu(\bar{x}_1), \mu(\bar{x}_2), \ldots, \mu(\bar{x}_n), F)$ to be inserted in R. Note that $V_0(D \cup \triangle D)$ remains unchanged. Since μ is a truth assignment, $\mu(x_i)$ and $\mu(\bar{x}_i)$ are complements for each $i \in [1, k]$. In other words, for each possible renaming δ of (X_i, Y_i, C) into X, Y, Z, $\pi_{K,X,Y,Z}(\delta(R))$ consists of tuples t of the form $(0, t[X, Y, Z])$ with $t[X, Y, Z]$ having a single T and two F-values. In other words, $V_1(D \cup \triangle D) = V \cup \triangle V$. In addition, there is exactly one T among the three literals of each clause C_j, and $Q_{2,j}$ takes all the permutations of the values of these three literals. Thus, again $V_2(D \cup \triangle D) = V \cup \triangle V$. That is, $Q(D \cup \triangle D) = V \cup \triangle V$ and hence $\triangle D$ is side-effect free. Furthermore, since no tuples could have been be added to R_v (due to the key constraints) and $\triangle D$ consists of a single tuple, $\triangle D$ is necessarily minimum.

Conversely, let $\triangle D$ be a minimum side-effect free solution such that $Q(D \cup \triangle D) = V \cup \triangle V$. Since $(0, F, F, T)$ is in $Q(D \cup \triangle D)$ and $\triangle D$ is side-effect free, we know that $\triangle D$ consists of a single tuple of the form $(0, a_1, a_2, \ldots, a_n, b_1, b_2, \ldots, b_n, F)$. Further, by the definition of V_1 and the fact that $Q(D \cup \triangle D) = V \cup \triangle V$, we have that the mapping $\mu(x_i) = a_i$ and $\mu(\bar{x}_i) = b_i$ is a truth assignment for ϕ. Finally, since $Q(D \cup \triangle D)$ is $\{(0, T, F, F), (0, F, T, F), (0, F, F, T)\}$, by the definition of V_2, μ

must be a 1-in-3 truth assignment for ϕ, since otherwise side effects would occur in $Q(D \cup \triangle D)$. □

Remark. From Proposition 5 and Theorem 6, it follows that any fragment of SPJU that contains both PU and J is both coNP-hard and NP-hard (lower bound), and is intractable. Hence, such a fragment is in a complexity class that subsumes NP and coNP (see, e.g., [36] for details about complexity hierarchies).

3.2.2 Tractability Results

The good news is that the problem is tractable for SP and SU views and for group insertions, no matter whether these views are key preserving or not.

Theorem 7. *The view side-effect problem is in PTIME for 1)* SP *views and 2) SU views, both for group insertions.*

Proof. The proof is constructive. For each of the cases, we provide a PTIME algorithm which either halts (indicating that no solution exists) or outputs a solution for the view side-effect problem. Note that some view insertions may not be doable, as shown in Example 3.1.

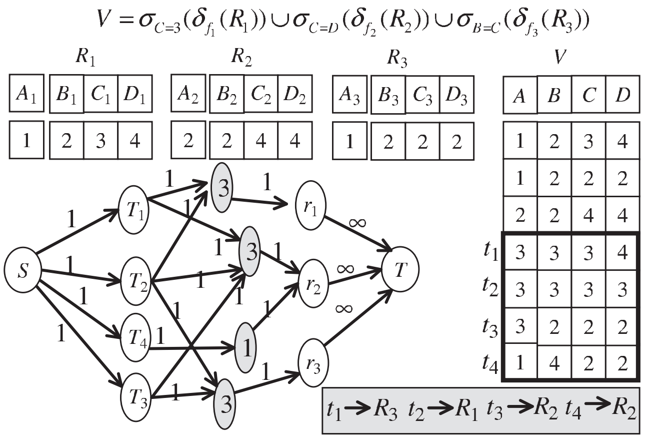
SP views. A PTIME algorithm can be found in the proof of [23, Theorem 4.3]. The algorithm first checks whether the tuples in group insertions satisfy the view definition, and whether the group insertions contain tuples with the same key attributes. Then for each view tuple, it generates its base tuples, fills in their missing values such that the selection conditions in the view definition are satisfied, and checks whether those tuples can be inserted into base tables without violating the keys.

Key preserving SU-views. Consider a key preserving SU query $Q = \bigcup_{i=1}^{k} Q_i$, an instance D of the schema $\mathcal{R}$, and a group insertion $\triangle V$. We assume that each S query Q_i is of form $Q_i = \sigma_{C_i}(\delta_{f_i}(R_{j_i}))$. Observe that if there exists a source insertion $\triangle D$ that produces $\triangle V$, then there exists a side-effect free solution. Indeed, since Q does not involve joins, one can always trim irrelevant tuples from the query result by adjusting values in $\triangle D$.

The PTIME algorithm first models the view side-effect problem as a flow network with integer capacity on each edge, and then computes the maximum flow of the network in polynomial time. If the value of the maximum flow equals to $|\triangle V|$, then we can get a solution to the view side-effect problem. Otherwise, there exists no solution for group insertions $\triangle V$.

We first model $\mathcal{R}$, Q, and $\triangle V$ as a flow network. The initial vertex set of the flow network is $\mathcal{N} = \{S, T\} \cup \{T_i \mid t_i \in \triangle V\} \cup \{r_i \mid R_i \in \mathcal{R}\}$, where S is the source node, T is the sink node, T_i represents a tuple node for each $t_i \in \triangle V$, and r_i represents a relation node for each relation R_i involved in V. The initial edge set consists of $\mathcal{E} = \{(S, T_i) \mid i \in [1, k]\} \cup \{(r_i, T) \mid i \in [1, n]\}$, where k is the size of $\triangle V$, n is the size of $\mathcal{R}$, the capacity of edges (S, T_i) is 1 and that of edge (r_i, T) is ∞.

We next encode $\triangle V$. For each tuple t in $\triangle V$, we check for each Q_i whether 1) t satisfies the selection condition C_i in Q_i; and 2) whether there exists no tuple in $Q_i(D)$ having conflicting key with t. If both conditions are satisfied for Q_i, the relation R_{j_i} in Q_i is called a *candidate host* for tuple t. If there exists no such a candidate host for t, the algorithm halts and no solution to the view side-effect problem exists. If all tuples in group update have at least one candidate host, then the algorithm continues.



$V = \sigma_{C=3}(\delta_{f_1}(R_1)) \cup \sigma_{C=D}(\delta_{f_2}(R_2)) \cup \sigma_{B=C}(\delta_{f_3}(R_3))$
R_1
R_2
R_3
V
S
T_1
T_2
T_4
T_3
3
3
1
3
r_1
r_2
r_3
T
1
∞
t_1
t_2
t_3
t_4
$t_1 \rightarrow R_3$ $t_2 \rightarrow R_1$ $t_3 \rightarrow R_2$ $t_4 \rightarrow R_2$

A_1	B_1	C_1	D_1
1	2	3	4

A_2	B_2	C_2	D_2
2	2	4	4

A_3	B_3	C_3	D_3
1	2	2	2

	A	B	C	D
	1	2	3	4
	1	2	2	2
	2	2	4	4
t_1	3	3	3	4
t_2	3	3	3	3
t_3	3	2	2	2
t_4	1	4	2	2

Fig. 4. Illustration of the algorithm of Theorem 7(2).

Note, however, that t will not necessarily be inserted into a candidate host. Indeed, whether t is inserted into one of its candidate hosts, say R_{j_i}, also depends on whether the other possible insertions (from $\triangle V$) into R_{j_i} cause a key violation together with the insertion incurred by t. The information regarding key information among tuples in their candidate hosts is modeled in the flow network as follows: for each candidate host R_{j_i} and each $t_\ell \in \triangle V$, we update the flow network $\mathcal{N}$. Let $\vec{a}_{j_i}$ be the tuple consisting of key attributes of t_ℓ in R_{j_i}. We distinguish between the following two cases:

- *Case 1.* There is no edge of the form $(v(\vec{a}_{j_i}), r_{j_i})$, where v is a vertex with $\vec{a}_{j_i}$ as its label. We add a new vertex v_{new} with $\vec{a}_{j_i}$ as label into $\mathcal{N}$, and add new edges (T_ℓ, v_{new}) and $(v_{\text{new}}, r_{j_i})$. The capacity of each new edge is 1, representing that at most one tuple with this key can be inserted in the host R_{j_i}.
- *Case 2.* There is already an edge of the form $(v(\vec{a}_{j_i}), r_{j_i})$, where v is a vertex with $\vec{a}_{j_i}$ as its label. This indicates that some other tuples (conflicting with t_ℓ on R_{j_i}) with the same key attributes $\vec{a}_{j_i}$ exist. Note that together with t_ℓ, only one such tuple can be inserted into R_{j_i}. Thus, we add an edge (T_ℓ, v) with capacity 1.

These steps complete the construction of flow network. The construction clearly runs in polynomial time. Below, we illustrate the construction. Fig. 4 shows the flow network for base relations $R_i(A_i, B_i, C_i, D_i)$, for $i = 1, 2, 3$, and SU view

$$V = \sigma_{C=3}(\delta_{f_1}(R_1)) \cup \sigma_{C=D}(\delta_{f_2}(R_2)) \cup \sigma_{B=C}(\delta_{f_3}(R_3)),$$

where δ_{f_i} renames A_i, B_i, C_i, D_i as A, B, C, D for $i = 1, 2, 3$. The group update $\triangle V$ is indicated by the bold rectangle. The gray shadowed vertexes are added following the rules in cases 1 and 2 and each of them has only one outgoing edge. As formally shown below, there is a solution for the insertion of $\triangle V$ into V iff the value of the maximum flow of the constructed flow network is $|\triangle V|$. For this example, there is a solution $\triangle D$ that inserts t_1, t_2, t_3, t_4 into R_3, R_1, R_2, R_2, respectively.

More formally, we show that there is a solution to the view side-effect problem iff the value of the maximum flow of the constructed flow network is $|\triangle V|$. First, we

TABLE 3
Complexity of the Source Side-Effect Problem for Deletion Propagation

Query class	single deletion		group deletion	
	general view	key preserving view	general view	key preserving view
JU, SJU, PJU, SPJU	NP-hard ([1])	NP-hard (Cor. 10)	NP-hard ([1])	NP-hard (Cor. 10)
SPU,SP,SU,PU	PTIME (Cor. 11)			
PJ, SPJ	NP-hard ([1])	PTIME (Prop. 8)	NP-hard ([1])	NP-hard (Cor. 9)
J, SJ	PTIME ([1])		NP-hard (Cor. 9)	

assume that the maximum flow $\phi : \mathcal{E} \to R^+$ equals $|\triangle V|$. By definition, this implies that $\sum_{(S,T_i)} \phi(S,T_i) = |\triangle V|$, or in other words, that ϕ assigns 1 to every edge starting from S. By the flow conversation law, i.e., for each $v \in \mathcal{N} \setminus \{S,T\}$, $\sum_{(v,w)\in\mathcal{E}} \phi(w,v) = \sum_{(v,w)\in\mathcal{E}} \phi(v,w)$, and the fact that all other edges (except those adjacent to T) have weight 1 assigned to them, this in turn implies that ϕ defines a unique path from S to T, one for each tuple in $\triangle V$. Since the in- and outgoing edges from the key-labeled vertexes have weight 1, these paths do not share key-labeled vertexes. Thus, from those paths one can infer the candidate hosts into which to insert each of the tuples. Indeed, for $t_i \in \triangle V$, one can simply follow the path (as defined by ϕ) starting form T_i until the corresponding host relation is reached.

Conversely, if a side-effect free solution exists, then there is a $\triangle D$ such that $Q(D \cup \triangle D) = V \cup \triangle V$. Since the view is key preserving, for any two distinct tuples $t_1, t_2 \in \triangle V$, there are different $t'_1, t'_2 \in \triangle D$ inserted into relations R_1, R_2, assuring that t_1 and t_2 appear in the view. Consider the following two cases: 1) R_1 and R_2 are different relations; and 2) R_1 and R_2 are the same relation R. For case 1, the construction of the flow network indicates that starting from S, we can reach T_1 and T_2 through different edges, and then reach different labeled nodes through different edges; finally, we reach different nodes r_1 and r_2. For case 2, the validity of the insertion guarantees that t_1 and t_2 have different keys on relation R. Thus, we can start from S and reach T_1 and T_2 via different edges, and then reach different labeled nodes, and finally reach node r. Since t_1, t_2 were chosen arbitrarily, this implies the existence of S to T paths passing through each tuple node in $\mathcal{N}$. Furthermore, all these paths share no labeled node and tuple node. This means that there is a feasible flow with value $|\triangle V|$.

Since for any flow ϕ, $\sum_{(S,T_i)} \phi(S,T_i) \leq |\triangle V|$, the maximum flow value is $|\triangle V|$. From this, it follows that the algorithm is in PTIME.

Arbitrary SU views. Since SU views do not have projections, they are key preserving. Hence, the PTIME algorithm above also works on arbitrary SU views. □

4 The Source Side-Effect Problem

In this section, we investigate the source side-effect problem. We study the problem for various key preserving SPJU views in Section 4.1 for both single-tuple and group deletions. Insertions are considered in Section 4.2.

4.1 Deletion Propagation

Given a view deletion $\triangle V$, the source side-effect problem for deletion propagation is to find a smallest set $\triangle D$ of source tuples to be deleted so that the tuples in $\triangle V$ are removed from the view.

Table 3 gives the complexity of determining the minimum $\triangle D$ for various subclasses of SPJU queries, for single-tuple and group deletions. Compared to Table 1, the results tell us that it is already hard to determine whether there exists $\triangle D$ to produce $\triangle V$, even without checking whether $\triangle D$ is side-effect free.

It has been shown that the source side-effect problem is already NP-hard for single deletions and PJ views [1]. We show that the problem for single deletions becomes polynomial-time solvable when the key preservation condition is imposed. This again verifies our observation that key preservation simplifies the analysis.

Proposition 8. *The source side-effect problem is in PTIME for key preserving* SPJ *views and single-tuple deletions.*

Proof. The PTIME algorithm presented in the proof of Proposition 1 is already able to compute a minimum source update $\triangle D$. We can therefore use the same algorithm for the source side-effect problem, except that we do not have to perform the steps for selecting the update that minimizes the number of view side effects. □

However, the problem for group deletions remains hard. Similar to Theorem 4, we show that the source side-effect problem is intractable for views defined with join only and for group deletions. This problem has not been considered by previous work.

Corollary 9. *The source side-effect problem is NP*-hard *for key preserving J views and group deletions.*

Proof. The proof of Theorem 4 suffices to show this. Indeed, the reduction given in that proof assures the minimality of the size of the source updates, and it does not impose any constraints on the size of side effects. □

For JU views, the problem is getting no easier, similar to its view side-effect counterpart (Corollary 2).

Corollary 10. *The source side-effect problem is NP-hard for key preserving JU views and single-tuple deletions.*

Proof. The proof of Corollary 2 is applicable here. Its reduction [1] concerns only the existence of a smallest source update that produces view updates. □

In contrast, for SPU views, the analysis is simpler, comparable to its view side-effect counterpart (Corollary 3).

Corollary 11. *The source side-effect problem is in PTIME for* SPU *views and group deletions.*

TABLE 4
Complexity of the Source Side-Effect Problem for Insertion Propagation

Query class	single insertion		group insertion	
	general view	key preserving view	general view	key preserving view
PJU, SPJU	NP-hard (Thm. 12)	PTIME (Cor. 17)	NP-hard (Thm. 12)	NP-hard (Thm. 13)
JU, SJU	PTIME (Cor. 16)		NP-hard (Thm. 13)	
PU, SPU			NP-hard (Thm. 14)	
PJ, SPJ	NP-hard (Thm. 12)	PTIME (Thm. 15)	NP-hard (Thm. 12)	PTIME (Thm. 15)
SU, SP, SJ	PTIME (Thm. 15)			

Proof. The PTIME algorithm given in the proof of Corollary 3 suffices to find smallest source updates. □

4.2 Insertion Propagation

Given a source database D, a query Q, the view $V = Q(D)$, and a set $\triangle V$ of tuples, the source side-effect problem for insertion propagation is to find a smallest set $\triangle D$ of tuples such that $Q(D \cup \triangle D)$ contains $\triangle V$, i.e., we want to find a smallest set of tuples to insert into the source database such that the insertion will get $\triangle V$ into the view, regardless of side effects on the view.

For single-tuple and group insertions, the complexity results for the source side-effect problem are summarized in the Table 4. Compared to its view side-effect counterpart (Table 2), the source side-effect problem is relatively easier for insertions since we no longer need to check whether source insertions are side-effect free.

4.2.1 Intractability Results

We first show that like its view side-effect counterpart (Proposition 5), the source side-effect problem is intractable for general PJ (and thus SPJ) views and single-tuple insertions. This tells us that the source side-effect analysis for insertions is more intriguing than its deletion counterpart (Proposition 8).

Theorem 12. *The source side-effect problem is NP*-hard *for PJ views and for single-tuple insertions.*

Proof Sketch. We prove this by reduction from the minimum set cover problem. The detail can be found in the proof of [23, Theorem 4.5]. □

The problem is no easier for *JU* views and group insertions, even when the views are key preserving.

Theorem 13. *The source side-effect problem is NP-hard for key preserving* JU *views and group insertions.*

Proof. We show this by reduction from the hitting-set problem. An instance of that problem consists of a collection C of subsets of a finite set S; it is to find a minimum subset X of S such that $X \cap c_i \neq \emptyset$ for all $c_i \in C$. The problem is NP-complete (cf. [35]).

Given S and C, we define an instance of the source side-effect problem. Let $S = \{x_i \mid i \in [1, n]\}$ and $C = \{c_j \mid j \in [1, k]\}$. We construct a source database D, a key preserving JU query Q, the view $V = Q(D)$, and a group insertion $\triangle V$. We show that we can find a minimum hitting set X of S iff there exists a minimum set $\triangle D$ such that $Q(D \cup \triangle D) \supseteq V \cup \triangle V$.

Base relations. For each $x_i \in S$ ($i \in [1, n]$), we define two relations $R_{(i,1)}(A_i, B_i)$ and $R_{(i,2)}(B_i, D_i)$ as follows:

- $R_{(i,1)}(A_i, B_i)$, where A_i is the key, and A_i, B_i range over $[1, k]$ and $\{T, F\}$, respectively. Intuitively, a tuple in $R_{(i,1)}$ indicates whether or not x_i belongs to a subset of C. Initially, (j, F) is in $R_{(i,1)}$ iff $x_i \notin c_j$.
- $R_{(i,2)}(B_i, D_i)$, where D_i is the key, and B_i, D_i range over $\{T, F\}$ and $[0, k]$, respectively. Intuitively, a tuple in $R_{(i,2)}$ indicates whether or not x_i is included in the hitting set. Initially, $R_{(i,2)}$ is empty.

View. We define a JU query $Q = \delta_{f_1}(Q_1) \cup \cdots \cup \delta_{f_n}(Q_n)$ where $Q_i = R_{(i,1)}(A_i, B_i) \bowtie R_{(i,2)}(B_i, D_i)$ and δ_{f_i} renames A_i, B_i, D_i as A, B, D, respectively. Initially, $V = Q(D) = \emptyset$. Intuitively, a tuple in the view is a triple (j, b, x), indicating whether at least one element of c_j is covered by the hitting set.

View insertions. The set $\triangle V$ is $\{(j, T, 0) \mid j \in [1, k]\}$.

The construction is illustrated in Fig. 5 for $S = \{a, b, c, d\}$ and $C = \{c_1 = \{a, b\}, c_2 = \{a, d\}, c_3 = \{b, c\}, c_4 = \{b, c, d\}\}$. The tuples inserted into the view and the tuples to be inserted into D are indicated by the bold rectangles. As will be seen shortly, $\triangle D$ in this example corresponds to a hitting set $X = \{a, b\}$ of C.

We next show the correctness of the reduction. First observe that for any source insertion $\triangle D$, if $V \cup \triangle V \subseteq Q(D \cup \triangle D)$, then $|\triangle D| \geq k + h$, where h is the size of minimum hitting set of C. To show this, we consider an index set $I = \{i \mid i \in [1, n]$, tuple $(T, 0)$ is inserted into $R_{(i,2)}$, $R_{(i,1)}$ accepts at least one tuple from $\triangle D\}$. Tuples inserted into $R_{(j,1)}$ or $R_{(j,2)}$ ($j \notin I$) are redundant, because they never generate any tuple in the view. Moreover, since $Q(D \cup \triangle D)$ contains $\triangle V$, for each $j \in [1, k]$, (j, T) must be inserted into some $R_{(i,1)}$ with $i \in I$. This means that each c_j must contain an element x_i with $i \in I$, which results in a hitting set $\{x_i \mid i \in I\}$ of C. Putting these together, $|\triangle D| \geq k + |I| \geq k + h$.

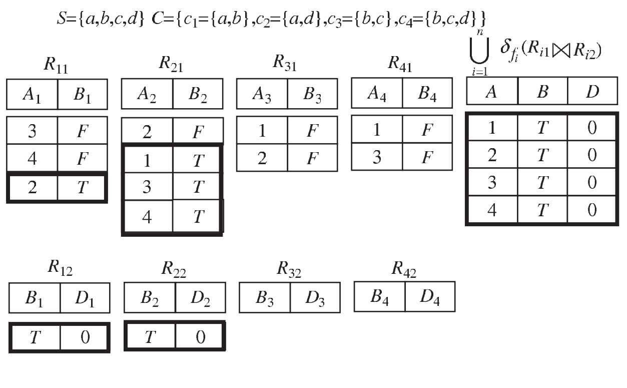


Fig. 5. Illustration of the proof of Theorem 13.

We next continue with verifying the correctness of the reduction. First, assume that X is a minimum hitting set of C. We then construct a set of source tuples $\triangle D$ such that $Q(D \cup \triangle D) \supseteq V \cup \triangle V$ and $|\triangle D|$ is minimum. Let $c_j \in C$ and take one x_i from $c_j \cap X$ that is not empty. Then, we insert (j, T) into relation $R_{(i,1)}$ and insert $(T, 0)$ into $R_{(i,2)}$. Note that by the keys, only one insertion can succeed. Hence, the total number of successful insertions of $(T, 0)$ into D will never exceed $|X| = h$. As a consequence, $(j, T, 0)$ belongs to $Q(D \cup \triangle D)$. Thus, $Q(D \cup \triangle D) \supseteq V \cup \triangle V$ and $|\triangle D| \leq k + h$. Together with the observation above, we see that $|\triangle D|$ is minimum.

Conversely, assume that $\triangle D$ is a minimum set such that $Q(D \cup \triangle D) \supseteq V \cup \triangle V$. We construct a minimum hitting set X of C. Let $I = \{i \mid \text{some tuple in } \triangle D \text{ is inserted into relation } R_{(i,1)}\}$. Since $|\triangle D|$ is minimum, there are no redundant insertions and thus $(T, 0)$ must be inserted into $R_{(i,2)}$ for all $i \in I$. Similarly, since $Q(D \cup \triangle D) \supseteq \triangle V$, we have that for each $j \in [1, k]$, (j, T) is inserted into D exactly once (into some relation $R_{(i,1)}$ with $i \in I$). Thus, $X = \{x_i \mid i \in I\}$ is a hitting set of C and $|\triangle D| = k + |I| = k + |X|$. If there exists another hitting set X' such that $|X'| < |X|$, we can use the construction method previously described to find a solution $\triangle D'$ such that $|\triangle D'| \leq k + |X'| < k + |X| = |\triangle D|$, which contradicts the assumption that $\triangle D$ is minimum. □

The problem for *PU* views is as hard as for *JU* views.

Theorem 14. *The source side-effect problem is NP-hard for key preserving* PU *views and group insertions.*

Proof. We prove this by reduction from the 1-in-3 3SAT problem. We refer to the proof of Theorem 6 for the statement of the 1-in-3 3SAT problem.

Given an instance ϕ of the 1-in-3 3SAT problem, we define a source database D, a key preserving PU query Q, and a set $\triangle V$ of tuples to be inserted into the view $V = Q(D)$ such that ϕ is 1-in-3 satisfiable iff there exists a (minimum) source insertion $\triangle D$ such that $V \cup \triangle V \subseteq Q(D \cup \triangle D)$. The reduction is similar to the one given in the proof of Theorem 6.

Base relations. The database D consists of one relation $R(K, X_1, \ldots, X_k, Y_1, \ldots, Y_k, C_1, \ldots, C_{n+k}, W)$, where K is the key, the attributes X_i and Y_i range over $\{T, F\}$ for $i \in [1, k]$, C_j ranges over $[1, n+k]$ for all $j \in [1, n+k]$, and W is in $\{F\}$. Intuitively, $X_i, Y_i (1 \leq i \leq k)$ encode the truth value of x_i and its negation, respectively. We use C_{n+i} together with W to determine whether a truth assignment is valid, i.e., whether only one of X_i and Y_i is T. We use C_j $(1 \leq j \leq n)$ to check whether clause c_j is satisfied by the truth assignment. Initially, D is empty.

View. We define a key preserving PU query $Q = V_1 \cup V_2$.

- $V_1 = \cup_{i=1}^{k} V_{1,i}$, where
$$V_{1,i} = \cup_{\delta \in \lambda(X_i, Y_i, W)} \pi_{K,X,Y,Z,C}(\delta_{n+i}(\delta(R))),$$
where $\lambda(X_i, Y_i, W)$ is the set of all (bijective) renaming of X_i, Y_i, W as X, Y, Z, and δ_{n+i} renames C_{n+i} as C. Intuitively, renaming in $\lambda(X_i, Y_i, W)$ enumerates permutations of X_i, Y_i, and W. The query V_1 is used to verify whether tuples in R define a valid truth assignment. Initially, $V_1(D) = \emptyset$.
- $V_2 = \bigcup_{i=1}^{n} V_{2,i}$, where each $V_{2,i}$ is defined according to the clause C_i of ϕ. Suppose that $C_i = \ell_{i_1} \vee \ell_{i_2} \vee \ell_{i_3}$. If $\ell_{i_j} = x_s$ for some $s \in [1, k]$, then let $A_{i_j} = X_s$; if $\ell_{i_j} = \bar{x}_s$, then let $A_{i_j} = Y_s$. We then define
$$V_{2,i} = \bigcup_{\delta \in \lambda(A_{i_1}, A_{i_2}, A_{i_3})} \pi_{K,X,Y,Z,C}(\delta_i(\delta(R))),$$
where as before, $\lambda(A_{i_1}, A_{i_2}, A_{i_3})$ is the set of all (bijective) renaming of $A_{i_1}, A_{i_2}, A_{i_3}$ as X, Y, Z, and δ_i renames C_i as C. Intuitively, $V_{2,i}$ introduces a tuple into V iff clause c_i is 1-in-3 satisfied by the truth assignment encoded by a tuple of D. Initially, $V_2(D) = \emptyset$.

View updates. Let $\triangle V = \{(0, \tau, j) \mid \tau = (\tau[X], \tau[Y], \tau[Z]), \tau \text{ is a permutation of } (T, F, F), j \in [1, n+k]\}$.

To show that this is a reduction, first assume that μ is a 1-in-3 truth assignment for ϕ. Let $\triangle D$ be the tuple
$$(0, \mu(x_1), \ldots, \mu(x_k), \bar{\mu}(x_1), \ldots, \bar{\mu}(x_k), 1, \ldots, n+k, F).$$
By the view definition and the fact that μ is a truth assignment, we have that $V_1(\triangle D)$ is the set $\{(0, \tau, j) \mid \tau = (\tau[X], \tau[Y], \tau[Z]) \text{ is a permutation of } (T, F, F), j \in [n+1, n+k]\}$. Similarly, we have that $V_2(\triangle D)$ is the set $\{(0, \tau, j) \mid \tau = (\tau[X], \tau[Y], \tau[Z]) \text{ is a permutation of } (T, F, F), j \in [1, n]\}$. Hence, $Q(D \cup \triangle D) = V \cup \triangle V$.

Conversely, assume that there exists a (minimum) source insertion $\triangle D$ such that $V \cup \triangle V = \triangle V \subseteq Q(\triangle D) = Q(D \cup \triangle D)$. Since the tuple $(0, T, F, F, 1) \in \triangle V$ is inserted into the view, R has key attribute K, and Q is key preserving, one can see that $\triangle D$ must contain a unique tuple t with $t[K] = 0$. Since all other tuples in $\triangle V$ share this key value, t is the only tuple in $\triangle D$ contributing to $\triangle V$. We can therefore assume that $\triangle D$ consists of the single tuple t. By the definition of V, each $V_{i,j}$ can only insert exactly one set of tuples into the view, where the set is of the form $\{(0, \tau, g(i,j)) \mid \tau \text{ is a permutation of } (T, F, F)\}$, and $g(i,j)$ is the value of $t[C_{n+j}]$ when $i = 1$ and it is $t[C_j]$ when $i = 2$. Since we know that there are precisely $n+k$ such sets in $\triangle V$, and there are precisely $n+k$ subqueries $V_{i,j}$, each mapping $g(i,j)$ must be a bijection. That is, $g : V_{i,j} \rightarrow [1, n+k]$ is a bijection, where $g(1,j) = t[C_{n+j}]$ and $g(2,j) = t[C_j]$.

We use the mapping g to define a truth assignment of $x_1, \ldots, x_k$. For arbitrary x_i, $(0, T, F, F, g(1,i))$ is inserted into V through a subquery $V_{1,i}$. By the definition of $V_{1,i}$, we can conclude that $\{t[X_i], t[Y_i]\} = \{T, F\}$. Hence, a truth assignment for ϕ is defined by $\mu(x_1) = t[X_1], \ldots, \mu(x_k) = t[X_k]$. Next, we verify that this assignment is a 1-in-3 truth assignment of ϕ. Consider an arbitrary clause C_i of ϕ. Since $(0, T, F, F, g(2,i))$ is inserted into V through subquery $V_{2,i}$, by the definition of $V_{2,i}$, we know that there is exactly one T among the truth values of the three literals in the clause. Hence, μ is indeed a 1-in-3 truth assignment. □

4.2.2 Tractability Results

We next identify some polynomial-time solvable subclasses. As in the view side-effect analysis (Theorem 7), *SP* views and *SU* views behave well. In contrast to their view side-effect counterparts (Proposition 5), the source side-effect analyses of *SJ* views and key preserving *SPJ* views also become simpler.

The result below also tells us that the source side-effect problem is in PTIME for SJ views and group insertions. In contrast, for group deletions, the problem is NP-hard for any views defined with join operations, no matter whether they are key preserving or not (Corollary 9).

Theorem 15. *The source side-effect problem is in PTIME for*

1. SP *views,*
2. SU *views,*
3. SJ *views, and*
4. *key preserving* SPJ *views,*

for group insertions.

Proof. The PTIME algorithms for cases 1 and 2 are similar to those of the view side-effect problem given in the proof of Theorem 7. In fact, the algorithms provided there return solutions for the source side-effect problem. We therefore concentrate on cases 3 and 4. These cases require a bit more effort (recall that the view side-effect problem for these cases is intractable, by Proposition 5).

SJ views. Consider an SJ query $Q = \sigma_C(\delta_{f_1}(R_1) \bowtie \cdots \bowtie \delta_{f_k}(R_k))$. A PTIME algorithm is given as follows:

Because Q does not contain any projection, we can derive for each tuple t in $\triangle V$ and for each relation R_i ($i \in [1, k]$) in Q a candidate insert tuple $\widehat{t_i} = (\vec{a}_i, \vec{b}_i)$ over the attributes of R_i, where $\vec{a}_i$ corresponds to the key attributes of R_i. We then check for each $t \in \triangle V$ whether $(\delta_{f_1}(\widehat{t_1}) \bowtie \cdots \bowtie \delta_{f_k}(\widehat{t_k}))$ satisfies the selection condition C. If not, then no solution exists and the algorithm halts. Otherwise, we check for each $\widehat{t_i}$ whether there exists already a tuple s_i in R_i having the same key $\vec{a}_i$. If this is the case, $\widehat{t_i}$ should be equal to s_i. If there exists a tuple $\widehat{t_i}$ for which this does not hold, then no solution exists and the algorithm stops. Otherwise, the algorithm continues.

Denote by $\triangle R_i$ the set of tuples $\widehat{t_i}$ for which no tuple in R_i exists with the same key. We finally check whether $\triangle R_i$ contains distinct tuples having the same key. If such tuples exists, no solution can be found and the algorithm halts. Otherwise, we define $\triangle D$ to be $\{\triangle R_1, \ldots, \triangle R_k\}$.

We remark that $\triangle D$ is the minimum solution. Indeed, in each instance R_i, the number of tuples inserted into D is the same as the number of new keys for R_i present in $\triangle V$. This is the minimum requirement for any solution.

The algorithm clearly runs in polynomial time.

Key preserving SPJ views. Consider a key preserving SPJ query $Q = \pi_{B_1,\ldots,B_m}\sigma_C(\delta_{f_1}(R_1) \bowtie \cdots \bowtie \delta_{f_k}(R_k))$.

As before, for each tuple t in $\triangle V$ and each relation R_i in Q, we associate a template $\hat{t}_i = (\vec{a}_i, \vec{b}_i, \vec{z}_i)$.

The PTIME algorithm first checks for incompatible templates. More specifically, the algorithm checks whether 1) there are no two different templates with the same key; and 2) no template $\hat{t}_i$ with the same key as an existing tuple s_i in R_i, but which differs from s_i in another attribute. If no incompatible templates are found, then the algorithm continues.

If no conflicts are found, we define $\triangle R_i$ to be the set of templates $\hat{t}_i$ which have no matching tuple s_i in R_i. We instantiate the variables in these templates as follows: for each tuple t in $\triangle V$, we compute a conjunctive formula ϕ_t representing the selection and join conditions to hold on $\hat{t}_1 \wedge \hat{t}_2 \wedge \cdots \wedge \hat{t}_k$, such that it will generate t in the view. The formula ϕ_t consists of conjuncts of equations of the form $x = y$, where x and y are variables or constants in $\hat{t}_i$, for $i \in [1, k]$. We group together all conjuncts ϕ_t into a single conjunction $\Phi = \wedge_{t \in \triangle V}\phi_t$, and check whether there exists an instantiation of the variables that satisfies Φ, using a method similar to the one given for case 1 in the proof of Theorem 7. An example of the algorithm can be found in the proof of [23, Theorem 4.4].

Since we are not concerned about the size of the side effects, we do not have to take into account constraints regarding existing constants in the database (this is in contrast to the coNP-hardness proof of Proposition 5). Hence, if an instantiation exists, we can convert the templates into tuples that populate the update set $\triangle R_i$. Finally, we define $\triangle D = \{\triangle R_1, \ldots, \triangle R_k\}$.

Obviously, $\triangle D$ is a solution. It is also minimum: at most a single tuple for each new key in tuples in $\triangle V$ is added, a necessary requirement for any solution. □

As opposed to Proposition 5 and Theorem 6, we show that the source side-effect problem is tractable for *SPU* and *SJU* views for single-tuple insertions. Putting these together with Theorems 13 and 14, we can see that group insertions complicate the source side-effect analysis.

Corollary 16. *The source side-effect problem is in PTIME for 1)* SPU *views and 2)* SJU *views, for single-tuple insertions.*

Proof. 1. *SPU* views. Consider an SPU query $Q = \bigcup_{i=1}^{k} \pi_{B_1,\ldots,B_m}(\sigma_{C_i}(\delta_{f_i}(R_i)))$, a database D, the view $Q(D)$, and view update $\triangle V$ consisting of a single tuple t. We present a PTIME algorithm that, given Q, D, $Q(D)$, and $\triangle V$, computes a minimum $\triangle D$ to produce $\triangle V$.

The algorithm first checks whether the only tuple t in $\triangle V$ is already present in the view $V = Q(D)$. If yes, then $\triangle D = \emptyset$. Otherwise, the algorithm proceeds as follows:

For $i \in [1, k]$, we invoke the PTIME algorithm for SP views given in the proof of Theorem 15 to check whether or not t can be inserted into $\pi_{B_1,\ldots,B_m}(\sigma_{C_i}(\delta_{f_i}(R_i)))$, by inserting a source tuple $\triangle D_i$ into base relation R_i. If the answer is yes for some $i \in [1, k]$, then it inserts a single tuple $\triangle D_i$ into the relation R_i. Otherwise, t cannot be inserted into V. Obviously, the algorithm is in PTIME, and $\triangle D$ contains at most one tuple (thus minimum).

2. *SJU* views. Similarly, one can develop a PTIME algorithm to handle SJU views and single-tuple insertions. The algorithm leverages the PTIME algorithm for SJ views (Theorem 15), along the same lines as above. □

Corollary 17. *The source side-effect problem is in PTIME for key preserving* SPJU *views, for single-tuple insertions.*

Proof. There exists a PTIME algorithm that, given a key preserving SPJU query Q, a database D, the view $Q(D)$, and view update $\triangle V$ with a single tuple as input, computes a minimum $\triangle D$ to produce $\triangle V$, if it exists. The algorithm is similar to the one for SPU views given in the proof of Corollary 16. Indeed, the only difference between the two is that here we invoke the PTIME algorithm for key preserving SPJ views (Theorem 15) for single insertions, rather than the one for SP views. □

5 The Annotation Placement Problem

In this section we investigate the annotation placement problem. Given a single location (field) l in a view tuple $\triangle V$, a source database D, an SPJU query Q, and the view $V = Q(D)$, the annotation placement problem is to identify a single field l' in a single tuple $\triangle D$ such that annotating l' is propagated to a smallest number of fields in the view including l. As stated in Section 1, this problem studies how an annotation attached to a location in the view is traced backward to the source data.

In contrast to the view side-effect problem and the source side-effect problem (Sections 3 and 4), we do not need to consider group updates or insertions for the annotation placement problem. As a consequence, this section presents a single result: the annotation placement problem is in PTIME for all subclasses of key preserving SPJU views. In contrast, it was shown in [1] that the problem is NP-*hard* for general PJ views, and it is in PTIME for SJU and SPU views.

Theorem 18. *The annotation placement problem is in PTIME for any subclass of key preserving* SPJU *views.*

Proof. It suffices to prove it for key preserving SPJU views. Let $\mathcal{R} = \{R_1, \ldots, R_n\}$ be a relational schema, D an instance of $\mathcal{R}$, and $Q = \cup_{i=1}^{k} Q_i$ an SPJU query, where $Q_i = \pi_{A_1,\ldots,A_m}(\sigma_{C_i}(\delta_{f_{i_1}}(R_{i_1}) \bowtie \cdots \bowtie \delta_{f_{i n_i}}(R_{i n_i})))$. We use (V, t, A_l) to denote a location in the view, i.e., an annotation is attached to the A_l attribute of tuple $t \in V$ $(= Q(D))$. We develop a PTIME algorithm that, given D, Q, $Q(D)$, and (V, t, A_l) as input, finds a single tuple $\triangle D$ in D and a single location $(D, \triangle D, B_l)$ in $\triangle D$ such that an annotation in this field of $\triangle D$ propagates to a smallest number of tuples in V including (V, t, A_l).

We first show how we process each SPJ subquery Q_i, $i \in [1, k]$. We decompose t into $t_{i_1}, \ldots, t_{i n_i}$ based on Q_i, where t_{i_j} is the attribute value vector (including key attributes) associated with the relation R_{i_j}. Utilizing the key preservation condition, the algorithm checks whether the following conditions are satisfied: 1) there is a tuple $t'_{i_j} \in R_{i_j}$ (unique if exists) with the same key-attribute values as t_{i_j} on R_{i_j} for each $j \in [1, n_i]$, 2) $t' = t'_{i_1} \bowtie \cdots \bowtie t'_{i n_i}$ satisfies the selection condition C_i, and 3) $t = \pi_{A_1,\ldots,A_m}(t')$. If not, we know that t cannot be generated by subquery Q_i and thus $\triangle D_i = \emptyset$. Otherwise, for each t_{i_j}, if it contains attribute B_l and an annotation on B_l can be propagated to the specified field A_l in t, we check the other tuples in the entire view $Q(D)$ to compute the side effects generated by annotating B_l of t_{i_j}. After processing all t_{i_j}, $j \in [1, n_i]$, we can find the location with the minimum side effects for annotating a tuple in $Q_i(D)$, as well as its side effects S_i on the entire view $Q(D)$, where S_i is the set of view locations affected by annotating B_l. This can be done in PTIME.

By processing each Q_i as above, we find $\triangle D_i$ and S_i for all $i \in [1, k]$. We pick a nonempty $\triangle D_j$ such that S_j is minimum among all S_i's with nonempty $\triangle D_i$. Then, $\triangle D_j$ is the location with minimum side effects. □

6 Conclusion

We have identified a practical condition, namely, the key preservation condition, which simplifies the propagation analysis of annotations. For key preserving views, we have shown that the annotation placement problem is tractable for all subclasses of SPJU queries, and that the view and source side-effect problems are in PTIME for SPJ views and single-tuple deletions, as opposed to NP-hard for general SPJ views [1], [12]. We have also investigated the impact of group updates on the complexity of the propagation analysis, and shown that group updates complicate the analysis: for group deletions, the view and source side-effect problems become NP-hard for all subclasses of key preserving SPJU views that involve join operation. In addition, we have established the first complexity results for the analysis of view insertions for SPJU views, key preserving or not. These provide a complete picture of the complexity (intractability and tractability) of the annotation propagation analysis, which is useful in both data provenance and view update processing.

We are currently studying approximation (heuristic) algorithms for the propagation analysis when the problems are intractable. We also plan to study the upper bounds for the intractable cases and to identify other practical conditions on view definitions such that the analysis can be performed efficiently. Finally, we only considered lower bounds for the intractable cases. We defer the study of upper bounds to future work.

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