

# Updating Recursive XML Views of Relations

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## Abstract

This paper investigates the view update problem for XML views published from relational data. We consider (possibly) recursively defined XML views, compressed into DAGs and stored in relations. We provide new techniques to efficiently support XML view updates specified in terms of XPath expressions with recursion and complex filters. The interaction between XPath recursion and DAG compression of XML views makes the analysis of XML view updates intriguing. Furthermore, many issues are still open even for relational view updates, and need to be explored. In response to these, on the XML side, we revise the update semantics to accommodate XML side effects based on the semantics of XML views, and present efficient algorithms to translate XML updates to relational view updates. On the relational side, we propose a mild condition on SPJ views, and show that under this condition the analysis of deletions on relational views becomes PTIME while the insertion analysis is NP-complete. We develop an efficient algorithm to process relational view deletions, and a heuristic algorithm to handle view insertions. Finally, we present an experimental study to verify the effectiveness of our techniques.

## 1 Introduction

Views provide an abstraction of the data stored in a database and are commonly used in practice. Commercial DBMSs have identified the need for materializing and/or providing ways of updating them, and propagating the updates to the underlying data [15, 22, 25]. Indeed, the study of relational views and their update mechanisms have received considerable attention (see, e.g., [10, 11, 16, 20]). Recently, a number of systems have been developed to publish relational data to XML [1, 4, 12, 15, 22, 25]; the published data are effectively XML *views* of the relational data. Thus, the problem of transparently updating the XML views needs to be revisited. Given an XML view of a relational database, we want to propagate updates of the XML view to the original *relational* tables, without compromising the integrity of neither the XML nor the relational data.

While several commercial systems [15, 22, 25] allow users to define XML views of relations, their support for XML view updates is either very restricted or not yet available. Previous work on XML view updates [2, 27, 28] has focused on translating XML view updates to relational view updates and delegating the problem to the relational DBMS; however, most commercial DBMSs only have limited view-

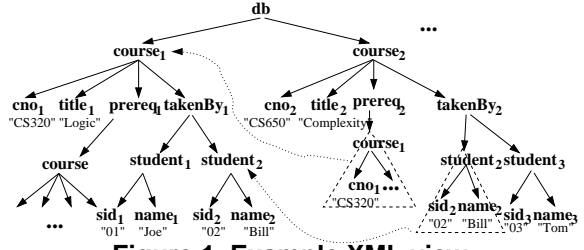


Figure 1. Example XML view

update capability [15, 22, 25]. The state of the art in XML view updates research explicitly focuses on *non*-recursively defined XML views and XML updates defined *without* recursive XPath queries. These restrictions are unfortunate since the recent proposals on XML update languages (e.g., [26]) employ recursive XPath queries while DTDs (and thus XML view definitions) found in practice are often recursive [6]. Given these requirements, we consider more general XML views and updates: possibly recursive XML view definitions and XML updates specified in terms of XPath expressions with recursion and complex filters, as illustrated below.

**Example 1.1:** Consider a *registrar* database with the following schema (keys are underlined): `course(cno, title, dept)`, `student(ssn, name)`, `enroll(ssn, cno)`, `prereq(cno1, cno2)`, where a tuple  $(c1, c2)$  in `prereq` indicates that  $c2$  is a prerequisite of  $c1$ . As depicted in Fig. 1 (the dotted lines will be explained shortly), an XML view  $T$  of the relational database is published for the CS department. The view is required to conform to the DTD below (the definition of elements whose type is PCDATA is omitted):

```
<!ELEMENT db      (course*)>
<!ELEMENT course (cno, title, prereq, takenBy)>
<!ELEMENT prereq (course*)>
<!ELEMENT takenBy (student*)>
<!ELEMENT student (ssn, name)>
```

The view is defined recursively since the DTD is recursive (`course` is indirectly defined in terms of itself via `prereq`). Consider an XML update  $\Delta_X$ , which inserts the subtree for `course` CS240, as a prerequisite of all courses given by the recursive XPath query `course[cno=CS650]//course[cno=CS320]/prereq`. To propagate  $\Delta_X$  means that we need to find an equivalent  $\Delta_R$  over the relational database that inserts the same information in the underlying tables so that if the database is re-published in XML it leads to the same XML view as the one we have after applying  $\Delta_X$  on  $T$ .  $\square$

The analysis becomes complicated since there are three sub-problems that cannot be treated in isolation, namely:

(i) how are the XML views efficiently materialized, (ii) what are the correct update semantics for XML views of relational data over the materialization primitives, and (iii) how are the new semantics implemented and the updates propagated to the materialized XML views *and* the relational database.

**Efficient materialization of XML views.** An XML document published from a relational database has high *compression* potential. In the document of Fig. 1 (Example 1.1), certain subtree instances can be shared; one can materialize each subtree shared by multiple nodes in the tree *only once*, as indicated in Fig. 1 (replacing the subtrees in the dotted triangles by dotted edges – e.g., the subtree for course CS320). The compressed view becomes a directed acyclic graph (DAG), which is often significantly (at times even exponentially) smaller than the original tree. Moreover, one may want to store the view (DAG) in *relations* itself. This is nontrivial: the XML view is recursively defined, and a naïve relational encoding may require an *unbounded* number of relational views. Further, the aim is to use recursive XPath expressions for denoting the parts of the document to be updated. Translation from (recursive) XPath queries over recursive XML views to SQL queries is hard [19]. To our knowledge, no efficient algorithm exists for evaluating XPath queries with *complex filters* on DAGs *stored in relations*. To this end, we present an efficient algorithm for evaluating XPath queries with *complex filters* on DAGs, based on a new and incrementally maintained indexing structure to handle recursion and a new technique for handling filters.

**XML update semantics.** Update semantics should be revised given the XML view materialization primitives. In Example 1.1, we are to insert CS240 as a *prereq* of *only* those CS320 nodes below CS650; however, CS320 nodes also occur elsewhere. As the XML view is published from the same relational database, all courses have unique *prereq* hierarchies. An insertion on *selected paths* of the hierarchy will result in *side effects* that should be detected. The users should then be consulted and, if they insist on continuing, the insert semantics needs to be revised so that the insertion will be performed at *every* CS320 node. The details of side effects on deletions are even more subtle and call for a new semantics. In light of this we refine the update semantics for XML views of relations to accommodate XML side effects. In addition, we develop an algorithm to translate *recursive* updates on a *possibly recursively defined* XML view to updates on the relational representation of the XML view.

**Update propagation.** Since the XML view is materialized in relations there is substantial work to be carried out in the relational realm. To this end, we identify a *key-preservation* condition on SPJ views, which is less restrictive than the conditions imposed by previous work [10, 11, 16]. We establish complexity results for the updatability problem to extend the few existing ones [3, 10]. We show that under

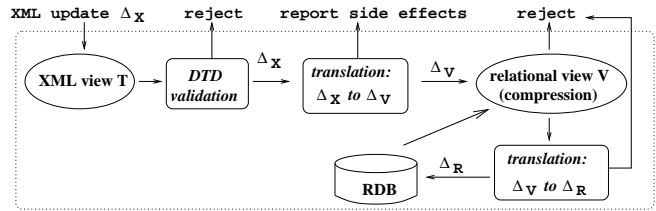


Figure 2. Overview of XML view updates

key-preservation on SPJ views, while the problem for tuple insertions is NP-complete, it becomes *tractable* for *group* deletions (which is NP-complete without key preservation). Given the complexity results, we propose a PTIME algorithm for processing group deletions on SPJ views and give an efficient heuristic algorithm to process group insertions.

**Problem statement and proposed solution.** Given an XML view defined as a mapping  $\sigma : \mathcal{R} \rightarrow D$  from relations of a schema  $\mathcal{R}$  to XML documents (trees) of a DTD  $D$ , a relational instance  $I$  of  $\mathcal{R}$ , the XML view  $T = \sigma(I)$ , and updates  $\Delta_X$  on the XML view  $T$ , we want to compute *relational updates*  $\Delta_R$  such that  $\Delta_X(T) = \sigma(\Delta_R(I))$ . That is, the relational updates  $\Delta_R$ , when propagated to XML via the mapping  $\sigma$ , yield the desired XML updates  $\Delta_X$  on the view  $T$ . We propose a framework for processing XML view updates, as shown in Fig. 2. For each XML view definition  $\sigma : \mathcal{R} \rightarrow D$ , we maintain a relational database  $I$  of  $\mathcal{R}$ , and the relational views  $V$  that encode the DAG compression of  $T = \sigma(I)$ . The users pose updates on  $T$  (Section 2). Given a single XML update  $\Delta_X$  on  $T$  as input, we generate a group update  $\Delta_R$  on  $I$  such that  $\Delta_X(T) = \sigma(\Delta_R(I))$  if such  $\Delta_R$  exists; otherwise *reject*  $\Delta_X$  as early as possible. Specifically, the framework processes an XML update  $\Delta_X$  on  $T$  in three phases, namely, *DTD validation* (see [7]), *translation from  $\Delta_X$  to  $\Delta_V$*  (Section 3), and *translation from  $\Delta_V$  to  $\Delta_R$*  (Section 4). If our algorithm detects a side effect, we report it to the user. After the relational update  $\Delta_R$  is computed, we update the underlying database  $I$  using  $\Delta_R$ , update the relational views  $V$  using  $\Delta_V$ , and finally, *in the background*, invoke our incremental algorithm to maintain our auxiliary structures. An experimental study is presented in Section 5, followed by related work in Section 6 and future work in Section 7. See the full version [7] for details.

## 2 View Updates Revisited in the XML Setting

We give a brief overview of publishing XML from relational data and present ways of efficiently materializing the XML view in relations. We then define the syntax and semantics of XML updates over this representation.

### 2.1 Schema-Directed XML View Definition

Our techniques are applicable to XML views published from relations via any system (e.g., Attribute Translation Grammars-ATG [1], SilkRoute, XPERANTO). We first briefly review ATG, a DTD-directed method for defining

XML views; we then present a way of materializing the published XML view in a relational database.

**DTDs.** A DTD  $D$  is a triplet  $(E, P, r)$ , where  $E$  is a finite set of *element types*;  $r \in E$  is called the *root type*;  $P$  defines the element types: for each  $A$  in  $E$ , there is a *production*  $A \rightarrow a$ , where  $a$  is a regular expression of the form:

$$\alpha ::= PCDATA \mid \epsilon \mid B_1, \dots, B_n \mid B_1 + \dots + B_n \mid B^*$$

where  $\epsilon$  is the empty word,  $B$  is a type in  $E$  (a *child type* of  $A$ ), and ‘+’, ‘,’ and ‘\*’ denote disjunction, concatenation and the Kleene star, respectively. A DTD is *recursive* if a type is defined (directly or indirectly) in terms of itself.

**XML views.** A publishing system implements a mapping  $\sigma : \mathcal{R} \rightarrow D$  from instances of a relational schema  $\mathcal{R}$  to documents of the target DTD  $D$ . (a) For each element type  $A$  of  $D$ ,  $\sigma$  defines a semantic attribute  $\$A$  whose value is a single relational tuple of a fixed arity and type; intuitively,  $\$A$  controls the generation of  $A$  elements in the XML view, and is used to pass data downwards as the document is produced. (b) For each production  $p = A \rightarrow \alpha$  in  $D$  and each type  $B$  in  $\alpha$ ,  $\sigma$  specifies a SPJ query,  $Q_{(A,B)}(\$A)$ , which extracts data from a relational database  $I$ , using  $\$A$  as a constant; it generates the  $B$  children of an  $A$  element and their  $\$B$  values. For example, for the production  $prereq \rightarrow course^*$ , the SPJ query  $Q_{prereq\_course}(\$prereq)$  can be specified as:

```

select  distinct c.cno, c.title
from    prereq p, course c
where   p.cno1 = $prereq and p.cno2 = c.cno

```

Intuitively, at a *prereq* node  $v$  with  $\$prereq$  value  $p$ , the subtree of  $v$  is constructed as follows: (1)  $Q_{prereq\_course}(p)$  is evaluated on the database  $I$ ; (2) for each distinct tuple  $c$  in the result of the query, a *course* child  $v_c$  of  $v$  is generated, which carries  $c$  as the value of its semantic attribute  $\$course$ ; and (3)  $c$  is then used in a similar fashion to expand the subtree rooted at  $v_c$ . The entire XML view is generated top-down starting from the root  $db$ , and conforms to the DTD of Example 1.1 (see [1, 7] for details).

**The subtree property and DAG compression.** An XML view of a relational database is determined by the underlying relational data. XML node uniqueness in this context is reflected as the *subtree property*. More specifically, consider a mapping  $\sigma : \mathcal{R} \rightarrow D$ . For any database  $I$  of  $\mathcal{R}$  and any type  $A$  of  $D$ , an  $A$ -element (subtree) in the XML view  $\sigma(I)$  is *uniquely determined* by the value of the semantic attribute  $\$A$  at its root. Thus, the publishing system in fact defines a function  $ST()$  such that, given an element type  $A$  and a value  $t$  of  $\$A$ ,  $ST(A, t)$  returns a subtree rooted at a node tagged  $A$  and carrying  $t$  as its attribute.

As noted in Section 1, a subtree  $ST(A, \$A)$  may appear at multiple places in the XML view  $\sigma(I)$ . It is natural and more efficient to *compress* the XML tree by storing a *single copy* of  $ST(A, \$A)$  no matter how many times it occurs in the XML view. This leads to a DAG representation of XML

view  $\sigma(I)$ . In Fig. 1, for example,  $course_1$  and  $student_2$  are shared subtrees (see dashed lines).

## 2.2 XML View Updates: Side Effects, Semantics

**Syntax.** We consider a class of XML updates [26] specified in terms of XPath: (a) insert  $(A, t)$  into  $p$ , (b) delete  $p$ . Here,  $A$  is an element type, and  $t$  is an instantiation of the semantic attribute  $\$A$  of  $A$ . Given the instantiation we can uniquely identify the root of a subtree of type  $A$ . We define  $p$  as an XPath expression ( $q$  in  $p[q]$  is called a *filter*):

$$\begin{aligned}
 p &::= \epsilon \mid A \mid * \mid // \mid p/p \mid p[q], \\
 q &::= p \mid p = 's' \mid label() = A \mid q \wedge q \mid q \vee q \mid \neg q,
 \end{aligned}$$

**Side effects.** On detecting side effects, users can choose either to abort the update, or to carry on under the semantics we provide in the sequence. Detection of side effects will be further elaborated in Section 3.2.

Recall that each subtree in the XML tree is *uniquely identified* by the value of the semantic attribute of its root. Moreover, under DAG compression, a single subtree may be shared among multiple parents. Therefore any changes to the subtree must be reflected to *all* instances of the subtree, *irrespective* of the XPath specified in the update operation. This forms the very basis for the appearance of side effects.

**Example 2.1:** In Example 1.1, a new subtree was to be inserted to change the prerequisite hierarchy of only those CS320 nodes below CS650. However, since there is a unique CS320 subtree, all changes to its prerequisite hierarchy must be reflected to *all* CS320 nodes, rather than only to those below CS650, leading to side effects.

Side effects are more subtle for deletions. Consider delete  $course[cno=CS650]/prereq/course[cno=CS320]$  on the same XML tree, that aims to remove course CS320 from the prerequisites of course CS650. This cannot be simply performed by physically removing all CS320 nodes as in previous work on XML view updates [2, 27, 28]: CS320 is itself an independent CS course and may be a prerequisite of other courses. For a correct deletion we need to find, for the root of the subtree to be removed, all its *parents* such that they are reachable via the XPath of the delete statement, *i.e.*, those *prereq* nodes ( $Prereq_2$ ) below CS650 nodes, and then remove CS320 from the *children* list of only those *parent* nodes. Note that CS320 is not removed from the *children* list of node DB since DB is not reachable via the XPath.  $\square$

**The semantics of XML view updates.** It is obvious that a new semantics should be developed to cope with *side effects*. This semantics needs to respect the hierarchical nature of XML views. Note that this semantics is *different* than the semantics of updates on XML data [26]. Given an XML view  $T$  with root  $r$ , an insert operation: (a) finds the set of all *elements* reachable from  $r$  via  $p$  in  $T$ , denoted by  $r[p]$ ; (b) for each element  $v$  in  $r[p]$ , it adds the new subtree

$ST(A, t)$  as the rightmost child of  $v$ ; and moreover, (c) for each element  $u$  that has the same type and semantic attribute value as  $v$ , it also adds  $ST(A, t)$  as the rightmost child of  $u$  as required by the semantics of XML views.

A deletion on XML views (a) computes  $r[[p]]$ ; (b) for each node  $v \in r[[p]]$ , it removes the subtree  $ST(A, t)$  from the children list of the parent node  $u$  of  $v$  such that  $u$  is reachable via XPath  $p$ , where  $A$  is the type of  $v$  and  $t$  is the value of  $\$A$  at  $v$ ; and (c) for any node  $u'$  of the *same type and semantic attribute value* as the parent  $u$  of  $v$ , it removes  $ST(A, t)$  from the children list of  $u'$ .

Compared to previous work [2, 27, 28], we support XML view updates that (a) are defined with much richer XPath expressions with *recursion and complex filters*, (b) operate on (possibly) recursively defined XML views, and (c) possess a new semantics that captures *side effects*, if any, of XML view updates. We also provide techniques to *detect* whether there are side effects and, in those cases, allow the users to cancel the update; otherwise, the operation will carry on with the semantics described earlier.

### 2.3 Relational Coding of Recursive XML Views

To reduce the update problem to a strictly relational one, we employ relational views to represent the XML views defined by a mapping  $\sigma : \mathcal{R} \rightarrow D$  from a relational schema  $\mathcal{R}$  to a DTD  $D$ . This is nontrivial: (a)  $\sigma$  is possibly recursively defined; on such views the encoding methods of previous work (e.g., [2]) may lead to *infinitely* many relational views; (b) we consider DAG compressions of XML views, i.e., a DAG representation of  $\sigma(I)$  where  $I$  is an instance of  $\mathcal{R}$  as opposed to *trees* assumed in previous work. To this end we define a relational representation  $\mathcal{V}_\sigma$  for the mapping  $\sigma$  by means of the edge relations in  $\sigma(I)$  as follows.

(a) We assume a compact, unique value associated with the tuple value of semantic attribute  $\$A$  in  $\sigma(I)$ . We assume w.l.o.g. the existence of a Skolem function [1]  $gen\_id$  that, given the tuple value of  $\$A$ , computes a unique  $id\_A$ . We use  $gen\_A$  to denote the set of the identities of all  $\$A$  tuples.

(b) We encode an XML view definition  $\sigma$  in terms of  $\mathcal{V}_\sigma$  as a set of SPJ queries  $Q_{edge\_A\_B}$  materializing the edge relations of  $\sigma$ . More specifically, for each production  $A \rightarrow P(A)$  in the DTD of  $\sigma$ , and for each child type  $B$  in  $P(A)$ , we create a relation  $edge\_A\_B$  with two columns,  $id\_A$  and  $id\_B$ . Consider productions of the form  $A \rightarrow B^*$ , where  $\$B \leftarrow Q(\$A)$  is the associated SPJ query in  $\sigma$ . Then  $edge\_A\_B$  is the set of pairs  $(ia, ib)$  such that  $ia = gen\_id(a)$ ,  $ib = gen\_id(b)$ , where  $a \in gen\_A$ ,  $b \in Q(a)$ . The definition of  $Q_{edge\_A\_B}$  is similar for productions of other forms. One example of an edge-relation query for the example of Fig. 1 is  $Q_{edge\_prereq\_course}$ :

```

select  gen_id(gp), gen_id(c.cno, c.title)
from    gen_prereq gp, prereq p, course c
where   p.cno1 = gp.cno and p.cno2 = c.cno

```

Observe the following about  $\mathcal{V}_\sigma$ . (1)  $\mathcal{V}_\sigma$  encodes the DAG *compression* of XML view  $\sigma(I)$ . Indeed, for any subtree  $ST(A, \$A)$  in  $\sigma(I)$ , each edge  $(ia, ib)$  in  $ST(A, \$A)$  is stored *only once* in a relation  $edge\_A\_B$  no matter how many times  $ST(A, \$A)$  (and thus the edge) appears in  $\sigma(I)$ . (2) Each  $Q_{edge\_A\_B}$  in  $\mathcal{V}_\sigma$  is defined by a SPJ query. Thus  $\mathcal{V}_\sigma$  consists of only SPJ views. (3)  $\mathcal{V}_\sigma$  consists of a *bounded* number of *relational views* even if  $\sigma$  is *recursively* defined. Indeed, each  $edge\_A\_B$  relation codes edges from  $A$ -nodes to  $B$ -nodes that may appear at an arbitrary depth of the tree, and the number of edge relations in  $\mathcal{V}_\sigma$  is *bounded* by the size of the DTD  $D$ .

**Updates on relational views.** Given an update  $\Delta_X$  on a DAG compressed XML view  $\sigma(I)$ , we convert it to updates  $\Delta_V$  on the relational view  $V = \mathcal{V}_\sigma(I)$ . The relational view updates  $\Delta_V$  consist of edge tuples of the form  $t = (ia, ib)$  to be inserted into or deleted from an edge relation  $edge\_A\_B$ .

To account for the side effects described earlier we compute the relational view updates  $\Delta_V$  such that (a) a newly inserted subtree is only stored once in  $V$  no matter how many times it appears in the updated view, and (b) a deleted subtree is not physically removed: only the tuple  $(ia, ib)$  in  $V$  representing the corresponding parent-child edge is deleted from its edge relation  $edge\_A\_B$ . More specifically, the tuple corresponding to  $ia$  is not removed from  $gen\_A$  because  $ia$  is a parent node in  $r[[p]]$  and needs to be kept in the XML view. To cope with subtree sharing,  $ib$  is not removed from  $gen\_B$  when the edge  $(ia, ib)$  is removed from  $edge\_A\_B$ ; instead, upon the completion of processing  $\Delta_V$ , our incremental maintenance algorithm runs in the *background* to remove tuples from  $gen\_B$ 's that are no longer linked to any node; at the completion of  $\Delta_V$   $gen\_B$ 's are updated (similarly for insertions).

## 3 Mapping XML View Updates to Relations

We present a technique for translating XML updates on an XML view to updates on relational views representing the DAG compression of the XML view. The technique consists of four parts: (a) indexing structures for checking ancestor-descendant relationships, (b) an efficient algorithm for evaluating XPath queries on DAGs and detecting side effects, (c) algorithms to translate updates on the XML view to updates on its relational representation, based on the indexing structures and the evaluation algorithm, and (d) incremental algorithms for maintaining the indexing structures.

### 3.1 Auxiliary Structures

To efficiently process recursion ('//') and filters in a DAG, we introduce two auxiliary structures: a *topological order* and a *reachability matrix*.

**Topological order.** Recall from Section 2 the function  $gen\_id()$ , which generates a unique id for each node based

<p><b>Input:</b> the relational view <math>V</math> and topological order <math>L</math>.</p> <p><b>Output:</b> reachability matrix <math>M</math>.</p> <ol style="list-style-type: none"> <li>1. <math>M := \emptyset</math>;</li> <li>2. <b>for</b> (<math>k :=  L </math>; <math>k &gt; 0</math>; <math>k--</math>) /*process <math>L</math> from right to left */</li> <li>3.     <math>d := L[k]</math>;</li> <li>4.     <math>A_d := \{a_2 \mid a_2 \in \text{anc}(a_1), a_1 \in \text{parent}(d)\}</math>;</li> <li>5.     insert <math>(a, d)</math> into <math>M</math> <b>for each</b> <math>a \in A_d</math>;</li> <li>6. <b>return</b> <math>M</math></li> </ol>
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**Figure 3. Algorithm Reach**

on the value of its semantic attribute. Given a representation of a DAG  $V$ , we create a list  $L$  consisting of all the distinct node ids in  $V$  topologically sorted such that  $u$  precedes  $v$  in  $L$  only if  $u$  is not an ancestor of  $v$  in the DAG, *i.e.*, there is no path from  $u$  to  $v$ . As will be seen shortly,  $L$  is useful in evaluating XPath filters as well as in computing and maintaining the reachability matrix. The list  $L$  can be computed in  $O(|V|)$  time (see, *e.g.*, [9]), where  $|V|$  is the size of the relational views. Its size,  $|L|$ , is the number of *distinct nodes* in the DAG, denoted by  $n$ . Note that  $L$  is computed once when  $V$  is created and it is maintained incrementally.

**Reachability matrix.** To efficiently evaluate ancestor-descendant relationship between pairs of nodes in a DAG, we use a conceptual *reachability matrix* encoded as a relation  $M(\text{anc}, \text{desc})$ , where  $\text{anc}$  is an ancestor node, and  $\text{desc}$  a descendant. We use  $\text{desc}(a)$  (resp.  $\text{anc}(a)$ ) to denote the descendants (resp. ancestors) of node  $a$  retrieved from  $M$ .

Relation  $M$  can be computed in  $O(|V|^2 \log |V|)$  time from  $V$  (see, *e.g.*, [9]). Capitalizing on the topological order  $L$  we give Algorithm Reach, shown in Fig. 3, that computes  $M$  in  $O(n |V|)$  time. It is based on dynamic programming: for a node  $d$ , the ancestors of the nodes in the set of parents of  $d$ , denoted by  $\text{parent}(d)$ , are already known before we compute ancestors  $A_d$ , such that we can compute  $A_d$  by using those previously computed ancestors (lines 4-5). Given the topological order guaranteed by  $L$ , this can be achieved by traversing  $L$  backwards (line 2). Note that  $\text{parent}(d)$  can be computed from the edge relations in  $V$ .

Algorithm Reach runs in  $O(n |V|)$  time: (a) for each node in  $L$  we visit its parents once and thus any node  $v$  is visited as many times as its in-degree, *i.e.*, the number of incoming edges to  $v$  in the DAG; (b) the sum of incoming edges to all nodes  $v$  is  $|V|$ ; (c) each visit takes at most  $O(n)$  time. In practice,  $|M| \ll n^2 \ll |V|^2$ , where  $V$  is even up to an exponential factor smaller than the XML tree  $T$ .

### 3.2 Evaluating XPath Queries on DAGs

To translate updates  $\Delta_X$  on XML views to updates  $\Delta_R$  on relational views and detect whether the update will yield side effects, we must evaluate the XPath expression used in  $\Delta_X$ . The DAG compression of XML views introduces new challenges: previous work on XPath evaluation has mostly focused on trees rather than DAGs. While evaluation algorithms were developed for path queries on DAGs [5, 23],

they cannot be applied in our setting because (a) they do not deal with complex filters which, as will be seen shortly, require a separate pass of the input DAG, and (b) they do not address maintenance of the indexing structures they employ, which is necessary when the DAG is updated. Path-query evaluation algorithms were also developed for semi-structured data (general graphs). However, these algorithms neither treat DAGs differently from cyclic graphs (and thus may not be efficient when dealing with DAGs), nor consider XPath queries used in XML view updates.

To this end we outline an efficient algorithm for evaluating an XPath query on an XML tree that is (a) compressed as a DAG, and (b) stored in edge relations  $V$ . The algorithm takes as input an XPath query  $p$  over XML tree  $T$ , the relational views  $V$ , and the reachability matrix  $M$ . It computes (a) a set  $r[[p]]$  consisting of, for each node reached by  $p$ , a pair  $(B, v)$ , where  $v$  is the id and  $B$  the type of the node respectively; (b) a set  $E_p(r)$  consisting of, for each  $v$  reached by  $p$ , tuples of the form  $((C, u), v)$ , where  $u$  is the id of a parent of  $v$  in the DAG such that  $p$  reaches  $v$  through  $u$ , and  $C$  is the type of  $u$ ; the set  $E_p(r)$  is needed for handling deletions; and (c) the set of nodes  $S$  in  $T$  which are affected by the update but are not reachable via  $p$ . If the set  $S$  is not empty, the update will generate XML side effects.

For XML data stored as a tree  $T$ , [18] developed an algorithm that evaluates an XPath query  $p$  in two passes of  $T$ . The basic idea of [18] is to first convert  $T$  to a binary-tree representation (before the two-pass process is invoked), and then run a bottom-up tree automaton on the binary tree to evaluate filters, followed by a run of a top-down tree automaton to identify nodes reached by  $p$ . It has linear-time complexity, the “optimal” one can expect [18]. We next show that a *comparable complexity* can be achieved when evaluating XPath queries on a DAG.

Our algorithm uses the following variables: (a) A list  $Q$  of filters including all the sub-expressions of filters in  $p$ , sorted such that for any  $q_i, q_j$  in  $Q$ ,  $q_i$  precedes  $q_j$  if  $q_i$  is a sub-expression of  $q_j$ . (b) For each  $q$  in  $Q$  and each node  $v$  in  $L$ , two Boolean variables  $\text{val}(q, v)$  and  $\text{desc}(q, v)$  to denote whether or not the filter  $q$  holds at  $v$  and at any descendant  $u$  of  $v$ , respectively. The algorithm has two phases: a bottom-up phase that evaluates *filters* in  $p$  and computes  $\text{val}(q, v)$  and  $\text{desc}(q, v)$  for each node  $v \in L$ , followed by a top-down phase that computes  $r[[p]]$  and  $E_p(r)$ . Due to lack of space we only outline the algorithm below.

**Bottom-up.** The key idea is based on dynamic programming. For each node  $v$  in the topological order  $L$ , and for each sub-filter  $q$  in the topological order  $Q$ , we compute the values of  $\text{val}(q, v)$  and  $\text{desc}(q, v)$ . This can be done by structural induction on the form of  $q$ . For example, when  $q$  is  $\text{label}() = A$ ,  $\text{val}(q, v)$  is true if and only if  $v$  is in  $\text{gen}_A$ . When  $q$  is  $q_1 \vee q_2$ ,  $\text{val}(q, v) := \text{val}(q_1, v) \vee \text{val}(q_2, v)$ . When  $q$  is a path expression  $p$ ,  $p$  can be rewritten

into a “normal form”  $\eta_1/\dots/\eta_n$ , where each  $\eta_i$  is either (a)  $\epsilon[q_i]$ , (b) a label  $A$ , (c) wildcard ‘\*’, or (d) ‘//’. The normal form can be obtained in  $O(|p|)$  time. Then, if  $q$  is rewritten as  $//\eta_2/\dots/\eta_n$  with  $\eta_1 = //$ ,  $\text{val}(q, v)$  is true if either  $\text{val}(\eta_2/\dots/\eta_n, v)$  or  $\text{desc}(\eta_2/\dots/\eta_n, u)$  is true for some child  $u$  of  $v$ ; correspondingly,  $\text{desc}(q, v)$  is true if either  $\text{val}(q, v)$  or  $\text{desc}(q, u)$  holds. The algorithm proceeds in the topological orders  $L$ . Thus the truth values of  $\text{val}(\eta_2/\dots/\eta_n, v)$  and  $\text{desc}(\eta_2/\dots/\eta_n, u)$  are already available before evaluating  $\text{val}(q, v)$  and  $\text{desc}(q, v)$ .

**Top-down.** We compute  $r[p]$ ,  $E_p(r)$  and  $S$  as follows. As mentioned,  $p$  can be rewritten as  $\eta_1/\dots/\eta_n$ , in which all the filters have already been evaluated to a truth value at each node. Starting from the root  $r$ , we find nodes  $C_i$  reached after each step  $\eta_i$  and maintain a set of nodes  $S$  in  $T$  that are not reachable via  $p$  but will be affected by the update. When  $\eta_i$  is ‘/’ (resp. ‘//’),  $S$  is extended with the parent (resp. ancestor) nodes of  $C_i$  that are not reached via  $p$ . These nodes can be found by using indexes on the edge relations  $V$  when  $\eta_i$  is  $A$  or \*, and by means of the reachability matrix  $M$  when  $\eta_i$  is ‘//’. The nodes reached by the last step  $\eta_n$  are put in  $r[p]$ , along with their types. The parents through which they are reached via  $p$  are put in  $E_p(r)$  along with their types. There is a side effect iff  $S$  is not empty. At that point, users may either abort the update, or continue using our update semantics.

**Complexity.** In the bottom-up phase, each node  $v$  is visited at most as many times as its incoming edges. In the top-down phase, each node is visited only once, except the final step when a node  $u$  may be included in  $E_p(r)$  at most as many times as its fan-out. The complexity of the algorithm is therefore  $O(|p| |V|)$ .

Observe the following: (a) When the DAG is a tree each node has one incoming edge and our algorithm visits each node at most twice, *i.e.*, it has the same complexity as that of [18]. When dealing with DAGs that do not have a tree structure, it is necessary to visit all the edges in a DAG in the worst case and thus our algorithm is optimal. (b) In contrast to [18], our algorithm does not require the conversion to binary trees and the construction of tree automata, which are potentially very large. (c) Our algorithm works on DAGs (including trees) while [18] cannot work on DAGs.

### 3.3 Translating Updates from XML to Relations

On account of the relational coding of XML views, a single XML update may be mapped to multiple relational updates (a group update) over the edge relations. We next give two algorithms, Xinsert and Xdelete, for translating XML view insertions and deletions to relational view updates.

**Insertion.** Algorithm Xinsert is presented in Fig. 4. Given  $\Delta_X = \text{insert}(A, t)$  into  $p$  on the XML view  $T$ , the algorithm returns the group insertions  $\Delta_V$  over  $V$  (which will then be tested for acceptance). We first compute the set of

**Input:** an insertion of the form  $\Delta_X = \text{insert}(A, t)$  into  $p$  over  $T$ , and the relational view  $V$ .

**Output:** a group insertion  $\Delta_V$  over  $V$ .

1.  $\Delta_V := \emptyset$ ;
2.  $E_A := \{ ((B, \text{gen\_id}(\$u)), (C, \text{gen\_id}(\$v))) \mid (u, v) \text{ is an edge in } \text{ST}(A, t), u, v \text{ with type } B, C \text{ resp.} \}$ ;
3.  $r_A :=$  the id of  $\text{ST}(A, t)$ ’s root as generated by  $\text{gen\_id}(t)$ ;
4. **for each**  $((B, ui), (C, vi)) \in E_A$
5.      $\Delta_V := \Delta_V \cup \{ \text{insert}(ui, vi) \text{ into } \text{edge\_}B\_C \}$ ;
6. **for each**  $(B, ui) \in r[p]$
7.      $\Delta_V := \Delta_V \cup \{ \text{insert}(ui, r_A) \text{ into } \text{edge\_}B\_A \}$ ;
8. **return**  $\Delta_V$ ;

**Figure 4. Algorithm Xinsert**

edges in the newly inserted subtree  $\text{ST}(A, t)$  rooted at  $r_A$ , according to the publishing mapping (lines 2-3), through function  $\text{gen\_id}()$ . We then generate the relational view updates: for each edge  $(ui, vi)$  in the newly inserted subtree, we add  $(ui, vi)$  to  $\Delta_V$  (lines 4-5); moreover, for each  $(B, ui) \in r[p]$ , we add  $(ui, r_A)$  as a new edge in  $\Delta_V$  (lines 6-7). The set  $r[p]$  of pairs  $(B, ui)$  of node identifiers along with their types reached by XPath  $p$  from the root of  $T$  (line 6) is computed using our XPath evaluation algorithm.

**Deletion.** Given  $\Delta_X = \text{delete } p$ , Algorithm Xdelete (not shown due to space constraints – see [7]) returns the group of deletions  $\Delta_V$  over the edge relations, which will be tested for acceptance (Section 4.2). For each node  $vi$  in  $r[p]$  and each parent  $ui$  of  $vi$  in  $E_p(r)$ , Xdelete removes the edge  $(ui, vi)$  from  $V$  (lines 2-3). The parent-child relation is computed by using the set  $E_p(r)$ , whose computation is coupled with that of  $r[p]$  (see Section 3.2).

**Example 3.1:** Consider the XML update  $\Delta_{X_1} = \text{delete } //\text{course}[cno=CS320]//\text{student}[sid=S02]$  on the XML tree in Fig. 1, which is to delete student S02 from the CS320 subtree. Given this as input, Algorithm Xdelete yields  $\Delta_{V_1} = \{(\text{takenBy}_1, \text{student}_2)\}$ .  $\square$

**Complexity.** Alg. Xinsert takes  $O(|E_A| + |r[p]|)$  time at most ( $|E_A|$  is the number of edges in  $\text{ST}(A, t)$ ). Alg. Xdelete takes  $O(|E_p(r)|)$  time. Added to  $O(|p| |V|)$  for evaluating  $p$ , this is the cost of generating  $\Delta_V$  from  $\Delta_X$ .

### 3.4 Maintaining Auxiliary Structures

The maintenance of auxiliary structures  $L$  and  $M$  is performed in the *background* in parallel with the processing of relational updates. What we ideally would like is to *incrementally* update  $M$ . Existing incremental techniques [14, 17] for updating reachability information are not applicable since they rely on special auxiliary structures which are themselves expensive to construct and maintain (*e.g.*, [14] requires the computation of a spanning tree, taking  $O(n |V|)$  time for each node insertion). Incremental algorithms of updating topologically ordered lists (*e.g.*, [21]) take  $O(|V|)$  time per edge insertion. We give a maintenance algorithm for  $M$  with  $O(n |V|)$  complexity by using  $L$ , and

**Input:** a deletion of the form  $\Delta_X = \text{delete } p$  over  $T$ , the rel. view  $V$ , reachability matrix  $M$  and topological order  $L$ .  
**Output:** deletions  $\Delta'_V$  over  $V$ ,  $\Delta_M$  over  $M$ , and updated list  $L$ .

1.  $\Delta'_V := \emptyset$ ;  $\Delta_M := \emptyset$ ;
2.  $L_R :=$  the sorted list  $\text{desc}(r[[p]])$  according to topological order  $L$ ;
3.  $\text{keep}(d) := \text{true}$  for each  $d \in T$ ; /\*initialize state \*/
4. **for each**  $d$  in  $L_R$  *traversed backwards*
5.      $P_d := \emptyset$ ;
6.     **for each**  $a \in \text{parent}(d)$
7.         **if**  $((C, a), d) \notin E_p(r)$  and  $\text{keep}(a) = \text{true}$
8.             **then**  $P_d := P_d \cup \{a\}$ ;
9.      $A_d := \{a_2 \mid a_2 \in \text{anc}(a_1), a_1 \in P_d\}$ ;
10.    **for each**  $a \in \text{anc}(d) \setminus A_d$
11.          $\Delta_M := \Delta_M \cup \{\text{delete}(a, d) \text{ from } M\}$ ;
12.    **if**  $P_d = \emptyset$  /\*compute  $\Delta'_V$  and update  $L$ \*/
13.         **then**  $\text{keep}(d) := \text{false}$ ;
14.             delete  $d$  from list  $L$ ;
15.             **for any child**  $d'$  (of type  $H$ ) of  $d$  (of type  $G$ )
16.                  $\Delta'_V := \Delta'_V \cup \{\text{delete}(d, d') \text{ from } \text{edge}_{G-H}\}$ ;
17. **return**  $(\Delta'_V, \Delta_M, L)$

**Figure 5. Maintenance algorithm  $\Delta_{(M,L)}\text{delete}$**

for  $L$  with  $O(n)$  time for each edge insertion using  $M$ .

Incremental maintenance in response to XML view deletions is given in Algorithm  $\Delta_{(M,L)}\text{delete}$  (Fig. 5). Due to space constraints we omit the maintenance algorithm  $\Delta_{(M,L)}\text{insert}$  for insertions (see [7]). The algorithm efficiently produces the following by scanning the elements of an XML deletion  $\Delta_X$ : (a) deletions  $\Delta_M$  over  $M$ , (b) an updated  $L$ , and (c) the set of edges  $\Delta'_V$  in the deleted subtree that are no longer connected to any nodes in the DAG and are to be passed to the garbage collector for *background* processing. The set  $\Delta'_V$  is a consequence of deletions  $\Delta_V$  computed by Algorithm  $X\text{delete}$ . The need arises when a node  $d \in \Delta_V$  is to be completely removed.

The algorithm progresses by populating deletions  $\Delta_M$  while, simultaneously, removing elements from  $L$  and populating  $\Delta'_V$ . The first step is arranging all nodes in all deleted subtrees in a list  $L_R$  (line 2): we compute  $\text{desc}(r[[p]])$ , *i.e.*, the descendants of all nodes in  $r[[p]]$ ; we then sort  $L_R$  according to  $L$ ; this is always possible since  $L_R \subseteq L$ . For each node  $d$  in the XML tree  $T$  we associate a state  $\text{keep}(d)$ , initialized to true, that keeps track of whether the node should be ultimately deleted or not (line 3).  $L_R$  is then traversed backwards (line 4); this processing order ensures that each  $d$  in  $L_R$  is processed after its ancestors thus guaranteeing correct deletion semantics. For each  $d$  in  $L_R$  we compute its undeleted parents (lines 6-8)  $P_d$  (*i.e.*, any node  $a$  in its parent set for which  $\text{keep}(a)$  is true) and then its *new* ancestors  $A_d$  (line 9). If there is a node in  $d$ 's current ancestors  $\text{anc}(d)$  that is not in  $A_d$ , it should be removed from  $M$  (lines 10-11). If  $d$  does not have any parents (*i.e.*,  $P_d = \emptyset$ ) we set  $\text{keep}(d)$  to false and delete it from  $L$  (lines 13-14). According to the semantics of  $L$ , an element removal does not affect the topological order. In addition, all outgoing edges from a deleted node  $d$  are deleted from  $V$  (lines 15-16); children  $d'$  of  $d$  can be readily identified from the edge relation determined by the types of  $d$  and  $d'$ .

**Example 3.2:** Recall  $\Delta_{X_1}$  of Example 3.1, Algorithm  $\Delta_{M,L}\text{delete}$  returns (1)  $\Delta'_V = \emptyset$ , (2) an unchanged  $L$ , and (3)  $\Delta_{M_1} = \{(\text{prereq}_2, \text{student}_2), (\text{prereq}_2, \text{sid}_2), (\text{prereq}_2, \text{name}_2), \dots\}$ , *i.e.*, the reachability information from nodes  $\text{prereq}_2$ ,  $\text{course}_1$  and  $\text{takenBy}_1$  to nodes in the S02 subtree ( $\text{student}_2$ ,  $\text{sid}_2$  and  $\text{name}_2$ ). Note that the set of edges  $\{(\text{takenBy}_2, \text{student}_2), (\text{takenBy}_2, \text{sid}_2), (\text{takenBy}_2, \text{name}_2), \dots\}$ , *i.e.*, the edges between  $\text{takenBy}_2$  (and thus  $\text{course}_2$ ) and the S02 subtree are still valid and are therefore not included in  $\Delta_{M_1}$ .  $\square$

**Complexity.** The worst-case time complexity of the algorithm is  $O(n |V|)$ , which is the cost of computing new ancestors for nodes in  $L_R$ . For each node in  $L_R$  we visit its parents once, which in total takes at most  $O(|V|)$  time (in practice it is much smaller than  $|V|$ ); at each visit, the algorithm takes at most  $O(n)$  time.

## 4 Updating Relational Views

We extend the study of relational view updates by providing complexity results and techniques for processing SPJ view updates under key preservation. All proofs are in [7].

### 4.1 Relational View Updates

**Key preservation.** Consider an SPJ query  $Q(R_1, \dots, R_k)$  that takes base relations  $R_1, \dots, R_k$  of  $\mathcal{R}$  as input, and returns tuples of the schema  $R(\vec{a})$ . We say that  $Q$  is *key preserving* if for each  $R_i$ , the primary key of  $R_i$  is included in  $\vec{a}$  (with possible renaming).

Key preservation is far less restrictive than other conditions proposed in earlier work for handling relational view updates (*e.g.*, [11, 16]). A mapping  $\sigma : \mathcal{R} \rightarrow D$  from a relational schema to a DTD employs SPJ queries [1]. Every SPJ query can be made key-preserving by extending its projection-attribute list to include the primary keys. Thus, in the sequel we assume w.l.o.g. that all the queries used in  $\sigma$  are key-preserving.

**Analysis.** Given a collection of views  $\mathcal{V}$  defined as SPJ queries *under key preservation*, a relational database  $I$  of schema  $\mathcal{R}$ , and a group view update  $\Delta_V$ , is there a group update  $\Delta_R$  on the database  $I$  such that  $\Delta_V(\mathcal{V}(I)) = \mathcal{V}(\Delta_R(I))$ ? In this setting,  $\Delta_V$  consists of either only tuple deletions or only tuple insertions, as produced by the translation algorithms of the last section. These deletions and insertions in  $\Delta_V$  are translated to deletions and insertions in  $\Delta_R$ , respectively. We use  $V$  to denote the view  $\mathcal{V}(I)$ . We refer to this problem as the *view updatability problem*.

It is known [3] that without key preservation, the updatability problem is NP-hard for a single deletion and a single PJ view, *i.e.*, when  $\Delta_V$  consists of a single deletion and  $\mathcal{V}$  is a view defined with projection and join operators only. We show that key preservation simplifies the updatability analysis for a collection of SPJ views and group deletions.

**Input:** a view definitions  $\mathcal{V}$ , a relational database  $I$ , the view  $V_Q = Q(I)$  for each  $Q \in \mathcal{V}$ , and a group deletion  $\Delta_V$ .

**Output:** a group update  $\Delta_R$  on  $I$  if it exists.

1.  $\Delta_R := \emptyset$ ;
2. **for each**  $(Q, t)$  in  $\Delta_V$
3.     compute  $Sr(Q, t)$ , the deletable source of  $t$  in  $V_Q$ ;
4. **for each**  $Q'$  in  $\mathcal{V}$  and **each**  $t$  in  $V_{Q'}$ , but not in  $\Delta_V$
5.     compute  $Sr(Q', t')$ ;
6. **for each**  $(Q, t)$  in  $\Delta_V$
7.     **if** there exists  $(S_j, t_j)$  in  $Sr(Q, t)$  such that  $(S_j, t_j)$  is not in  $Sr(Q', t')$  for any  $Q'$  in  $\mathcal{V}$  and any  $t'$  in  $V_{Q'}$ , but not in  $\Delta_V$
8.     **then**  $\Delta_R := \Delta_R \cup \{(S_j, t_j)\}$ ;
9.     **else** reject  $\Delta_V$  and **exit**;
10. **return**  $\Delta_R$

**Figure 6. Algorithm delete**

**Theorem 4.1:** For group view deletions  $\Delta_V$ , the SPJ view updatability problem is in PTIME.  $\square$

The problem is intractable for insertions under key preservation; the lower bound is verified by reduction from the non-tautology problem, which is NP-complete (cf. [13]).

**Theorem 4.2:** The SPJ view updatability problem is NP-complete even when  $\Delta_V$  has a single insertion and  $\mathcal{V}$  has a single view.  $\square$

These are the *first* complexity results for relational view updates under key preservation.

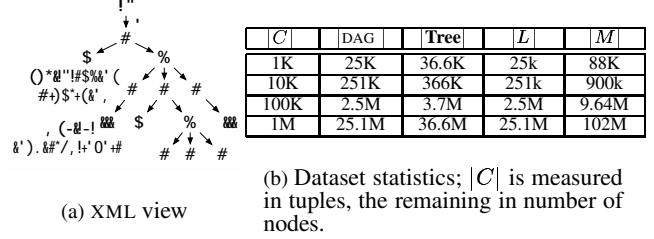
## 4.2 Handling Group Updates

We give a PTIME algorithm for computing database tuple deletions  $\Delta_R$  from a group of view deletions  $\Delta_V$ .

Let  $V_Q$  be the view  $Q(I)$ , and consider a tuple  $t$  in  $\Delta_V$  to be deleted from  $V_Q$ . Key preservation allows us to identify, for each relation  $S_j$  of  $Q$ , a *unique* tuple  $t_j$  via its key in  $t$ , such that  $t_1, \dots, t_i$  produce  $t$  via  $Q$ . Let  $Sr(Q, t)$  denote the set consisting of all the pairs  $(S_j, t_j)$ , referred to as the *deletable source* of  $t$  in  $V_Q$ . Observe the following. (a) Deleting any  $t_j$  from  $S_j$  suffices to remove  $t$  from  $V_Q$ . (b) Deletion of a source tuple  $t_j$  from  $V_Q$  is *side effect free* if and only if  $(S_j, t_j)$  is not in the deletable source of any tuple  $t' \in \mathcal{V}(I) \setminus \Delta_V$  that is to remain in the view after  $\Delta_V$  is performed. Hence,  $t$  can be deleted from  $V_Q$  if and only if there exists  $(S_j, t_j) \in Sr(Q, t)$  such that for all  $Q' \in \mathcal{V}$  and all  $t'$  that are in  $Q'(I)$  but not in  $\Delta_V$ ,  $(S_j, t_j)$  is not in  $Sr(Q', t')$ . Note that when *only the updatability problem* is concerned, deleting any of such  $t_j$  suffices.

Based on this we give Algorithm delete in Fig. 6, which is self-explanatory. The worst-case complexity of Algorithm delete is in  $O(|\Delta_V|(|\mathcal{V}(I)| - |\Delta_V|))$  time (see [7]).

**Minimal deletions.** Algorithm delete does not address which  $\Delta_R$  to select if multiple valid  $\Delta_R$ 's exist. In such cases, it is natural to choose the *smallest* set  $\Delta_R$  of tuples to delete, *i.e.*, a set  $\Delta_R$  such that  $|\Delta_R|$  is minimized. The *minimal view deletion problem* is thus to find, given a collection  $\mathcal{V}$  of view definitions, a database  $I$  and view deletions  $\Delta_V$ , the smallest set of tuple deletions  $\Delta_R$  such that  $\Delta_V(\mathcal{V}(I)) = \mathcal{V}(\Delta_R(I))$ . However desirable, the minimal



**Figure 7. Description of the datasets**

view deletion problem is intractable, even under key preservation. The lower bound can be verified by reduction from the NP-complete minimal set cover problem (cf. [13]).

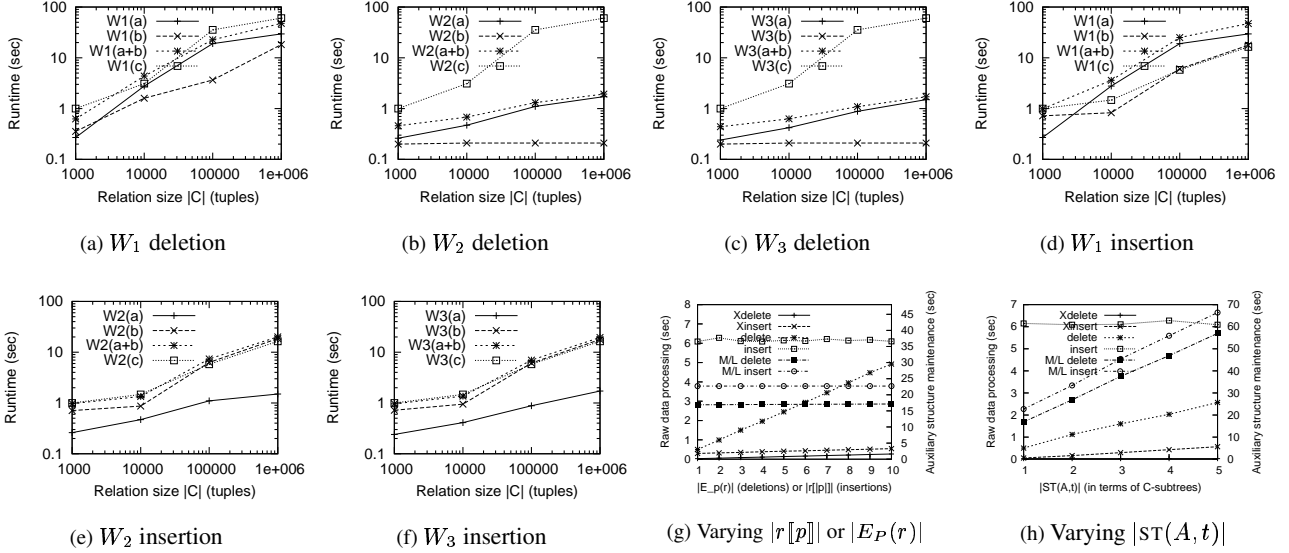
**Theorem 4.3:** For SPJ views under key preservation, the minimal view deletion problem is NP-complete.  $\square$

**Group Insertions.** Theorem 4.2 tells us that any practical algorithm for handling group view insertions is necessarily heuristic. We approach this by reducing the SPJ view insertion problem to SAT, one of the most studied NP-complete problems. This allows us to leverage a well-developed SAT solver [24] to efficiently compute  $\Delta_R$  if it exists. Due to space constraint the algorithm is omitted but is given in [7].

## 5 Experimental Study

We conducted a preliminary experimental study of our proposed view update mechanism in order to verify its effectiveness. The reported numbers are warm numbers and are the average of five runs per query. The standard deviation of the reported numbers is 5%.

All experiments were conducted on a dataset of four base relations:  $C(\underline{c}_1, \dots, c_{16})$ ,  $F(\underline{f}_1, \dots, f_{16})$ ,  $H(\underline{h}_1, \underline{h}_2)$  and  $C_U(\underline{c}'_1, \dots, c'_{16})$ , where underlined attributes indicate keys. The domain of  $f_1$  was equal to that of  $c_1$  and  $c'_1$ . The remaining  $C$  and  $F$  attributes controlled how many joining  $C$  and  $F$  tuples were filtered out. The domains of  $h_1$  and  $h_2$  were the same as that of  $c_1$ . In addition (1) for each  $c \in C \cup C_U$  there would be on average three tuples  $h \in H$ , where  $c_1 = h_1$ , and (2)  $h_1 < h_2$ , where  $(h_1, h_2) \in H$ . The universe of  $C$ , namely  $C_U$ , consisting of 100M  $C$ -tuples, ensured that whenever  $h_2$  joined with  $c_1$  it always yielded a  $C$ -tuple. The sizes of  $F$  and  $H$  were proportional to the size of  $C$ , used for reporting the size of the database; specifically, we report  $|C|$ , which ranges from 1,000 to 1,000,000 tuples, while  $|F| = |C|$  and  $|H| \simeq 3|C|$ . We defined an XML view of  $C, F$  and  $H$ ; as indicated in Fig. 7(a), the  $C$  nodes in the view were recursively defined, and a recursion of  $C$  in the view can be understood as  $\pi_{c_1, f_1, h_1, h_2}(\sigma_{c_1=f_1 \wedge f_1=h_1 \wedge h_2=c'_1 \wedge c_2=f_2 \wedge c_3=f_3 \wedge c_4=f_4} (C \times F \times H \times C_U))$ . Here  $C$  subtrees are shared, and subtree sharing accounted for 31.4% of  $C$  instances. Figure 7(b) lists some statistics on the number of published  $C$  subtrees and their compressed DAGs, and the corresponding sizes of the reachability matrix  $M$  and topological order  $L$ .



**Figure 8. Update performance as a function of the sizes of the relational database and the view update**

**Varying database size.** We generated two random update workloads over the XML view, one for insertions, and one for deletions; each workload consisted of three update classes, each class including ten operations. The classes were characterized by the XPath queries used to define the updates. Class  $W_1$  used XPath queries using the descendant axis and value filters; XPath queries in  $W_2$  used the child axis and value filters; finally,  $W_3$  contained XPath queries using the child axis and both structural and value filters. The times we report include: (a) the time to evaluate XPath queries; (b) the time to translate  $\Delta_X$  to  $\Delta_V$  (Algorithms Xinsert and Xdelete) and subsequently  $\Delta_V$  to  $\Delta_R$ , and the time to execute the update; and (c) the time to maintain the auxiliary structures in the *background* (Algorithms  $\Delta_{(M,L)}$ insert and  $\Delta_{(M,L)}$ delete).

Figures 8(a), 8(b) and 8(c) show the performance of the deletion algorithms for  $W_1$ ,  $W_2$  and  $W_3$ , respectively. We plot the runtime of performing the updates broken into their (a), (b) and (c) above constituents for various database sizes. Note that both axes use a logarithmic scale. The algorithms scale linearly with the size of the relational database. As shown, deletion time is dominated by XPath evaluation. Although the cost for auxiliary structure maintenance is relatively high, it is performed in the background.  $W_1(b)$  is the highest reported time among the three workloads since its XPath queries generate more edges (*i.e.*, a greater  $|E_p(r)|$ ), which are then examined by Algorithm delete.

Similar results are reported for insertions, as shown in Figures 8(d), 8(e) and 8(f) for  $W_1$ ,  $W_2$  and  $W_3$ , respectively (again, using logarithmic scales). The size of the inserted subtree was fixed. The SAT solver [24] returned a truth assignment in 78% of the cases and we only report the time for insertions where the SAT solver successfully returned a truth assignment. As for deletions, our insertion algorithms scale linearly with the size of the relational database.

**Varying update size.** For these experiments, we fixed  $|C|$  to 100K tuples. Figure 8(g) shows the performance of each algorithm as we varied  $|E_p(r)|$  (see Section 3.2) for deletions and  $|r[p]|$  for insertions, while keeping  $|ST(A,t)|$  constant to a single  $C$ -subtree. The runtimes for Algorithms Xinsert, Xdelete, delete and insert are shown on the left  $y$ -axis and the runtimes for algorithms  $\Delta_{(M,L)}$ insert and  $\Delta_{(M,L)}$ delete are shown on the right one. The translation time from  $\Delta_X$  to  $\Delta_V$  for Algorithm Xinsert (resp. Algorithm Xdelete) increases slightly as  $|r[p]|$  (resp.  $|E_p(r)|$ ) increases. The slope for Algorithm delete is large, as the increase of  $|E_p(r)|$  involves more queries to determine the source tuples to be deleted. The performance of Algorithm insert is dominated by the coding time. As  $|C|$  is far larger than  $|ST(A,t)|$  and  $|r[p]|$ , and the number of database queries required remains fixed, the coding time remains roughly constant though the size of the resulting coding increases; that only results in a non-observable increase in the SAT solver's runtime keeping the curve relatively flat. The performance of Algorithm  $\Delta_{(M,L)}$ insert (which can be found in [7]) and Algorithm  $\Delta_{(M,L)}$ delete is almost unaffected by  $|r[p]|$  (resp.  $|E_p(r)|$ ) since  $|ST(A,t)|$  is fixed.

Similar results are shown in Fig. 8(h) where we varied the size of  $|ST(A,t)|$  while fixing  $|E_p(r)| = 1$  and  $|r[p]| = 1$ . The performance of Algorithm Xdelete remains unchanged and its runtime is negligible for a fixed  $|E_p(r)|$ . Algorithm Xinsert scales linearly with the update size  $|ST(A,t)|$  as it needs to process  $ST(A,t)$  to generate  $\Delta_V$ . Algorithms  $\Delta_{(M,L)}$ insert and  $\Delta_{(M,L)}$ delete evidently scale linearly with the update size for similar reasons.

**Effectiveness of incremental maintenance.** The cost of incrementally maintaining the reachability matrix  $M$  and the topological order  $L$  is shown in Table 1. The first column is the size of the database. The total time needed for incrementally maintaining both auxiliary structures is given

Sizes	Incremental (Sec.)		Recomputation (Sec.)	
	$C$	Insertion	Deletion	$L$
1K	1.0	1.0	6.3	9.8
10K	4.6	3.1	86	288
100K	22.7	16.9	631	3,600
1M	84.2	61.5	8611	14,000

**Table 1. Incremental maintenance of  $L$  and  $M$**

in the second column for Algorithm  $\Delta_{(M,L)}$  insert and in the third column for Algorithm  $\Delta_{(M,L)}$  delete. The time for recomputing each structure is shown in the last two columns. The advantages of incremental maintenance become more prominent as the size of the data increases.

## 6 Related Work

Commercial database systems [15, 22, 25] provide support for defining XML views of relations and restricted view updates. IBM DB2 XML Extender [15] supports only propagation of updates from relations to XML but not vice-versa. Oracle XML DB [22] does not directly allow updates on XML (XMLType) views. In SQL Server [25], users specify the “before” and “after” XML views using *updategram* instead of update statements; the system then computes the difference and generates SQL update statements. The views supported are very restricted: only key-foreign key joins are allowed; neither recursive views nor updates defined in terms of recursive XPath expressions are supported.

There have been recent studies on updating XML views published from relational data [2, 28]. In [2], XML views are defined as *query trees* and are mapped to relational views. XML view updates are propagated to relations only if XML views are well-nested (*i.e.*, key-foreign key joins), and if the query tree is restricted to avoid duplication. An analysis on deciding whether or not an update on XML views is translatable to relational updates, along with detection algorithms, are provided in [28] and demonstrated in [27]. The limitations of previous work [2, 27, 28] are discussed in Section 1.

There has been a host of work ([10, 11, 15, 16, 20, 22, 25]) on relational view updates. [11] provides algorithms for handling restricted view updates without side effects in the presence of functional dependencies. The algorithm in [16] studies updates (with side effects) on a restricted class of SPJ view: key-foreign key joins and join attributes must be preserved. Our key preservation condition is less restrictive than that of [11, 16]. There has also been work ([10, 20]) on relational view complements, an approach quite different from ours. Commercial DBMSs [15, 22, 25] allow updates on very restricted views.

Few complexity bounds are known for (relational) view updates. The complexity of deletion on views is given in [3]. To our knowledge, our work is the first to establish complexity bounds for both deletion and insertion on views under key preservation.

A number of XPath evaluation algorithms have been proposed (*e.g.*, [8, 18]) for trees and cannot evaluate XPath

queries on DAGs. Path query evaluation for DAGs has been studied in [5, 23]; these approaches cannot be directly used in the context of XML view updates (see Section 3.2).

## 7 Conclusions

We have proposed new techniques for updating XML views published from relational data. We plan to extend our techniques to handle more general XML updates proposed in [26], and to find minimal, side-effect-free relational updates in response to XML view updates.

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