Additive First-Order Queries

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Problem statement

Characterize the FO queries that can be answered by only looking at domain-disjoint subinstances

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Additive First-Order Queries
A fact has the form $R(a_1, \ldots, a_k)$ where $R/k$ is a relation name and $a_1, \ldots, a_k$ come from an infinite universe of data elements.

A database schema $S$ is a finite set of relation names.

An instance $I$ is a nonempty set of facts over $S$.

The active domain $\text{adom}(I)$ is set of all data elements in $I$.

A $k$-ary query over $S$ maps each instance $I$ over $S$ to a $k$-ary subset on $\text{adom}(I)$.
A query $Q$ is additive if $Q(I \cup J) = Q(I) \cup Q(J)$ for any instances $I, J$ such that $\text{adom}(I) \cap \text{adom}(J) = \emptyset$

We denote the class of additive queries with ADD

Interesting class of queries:
- Allows coordination-free distributed query evaluation
- Useful in the analysis of the expressiveness of query languages

Examples:
- Computing the union and difference of relations are additive
- Computing the cartesian product is **NOT** additive
An additive query

\( Q = \) “Select the nodes with degree at least 2”

\[
Q = \{\{\ldots\}\} = Q \bigcup Q
\]

Expressible in FO: \( \{(x) \mid \exists y \exists z((R(x, y) \land R(x, z)) \land y \neq z)\} \)
CFO is a syntactical fragment of FO:

- Conjunction $\varphi \land \psi$ is only allowed when $\text{free}(\varphi) \cap \text{free}(\psi) \neq \emptyset$
- Disjunction $\varphi \lor \psi$ is only allowed when $\text{free}(\varphi) = \text{free}(\psi)$
- Negation is only allowed in the form $\varphi \land \neg \psi$ where $\text{free}(\psi) \neq \emptyset$, and $\text{free}(\psi) \subseteq \text{free}(\varphi)$
- For all quantification abbreviations are not allowed

Example: $\varphi(x) = \exists y \exists z ((R(x, y) \land R(x, z)) \land \neg (y = z))$ is in CFO
Theorem

$FO \cap ADD = CFO$

$\subseteq$: CFO is additive

$\subseteq$: Different proofs for:
  - Formula with at least one free variable
  - Sentences

The theorem holds over all instances, and over finite instances
Every additive FO formula $\varphi$ is $2^{qr(\varphi)}$-local

A formula $\varphi(\bar{x})$ is called $\ell$-local if for every instance $I$ and every tuple $\bar{a}$:

$$\bar{a} \in \varphi(I) \iff \bar{a} \in \varphi(N^I(\bar{a}, \ell))$$

where $N^A(\bar{a}, \ell)$ is the restriction of $I$ to the ball $B(\bar{a}, \ell)$.
Every additive FO formula \( \varphi \) is \( 2^{\text{qr}(\varphi)} \)-local

Let \( \varphi(x_1, \ldots, x_n) \) be an additive formula. For every \((a_1, \ldots, a_n) \in \varphi(I)\) we have \( d^I(a_i, a_j) \leq (n - 1)2^{\text{qr}(\varphi)} \).

Both Lemmas are proven using an Ehrenfeucht-Fraïssé game argument inspired by Otto’s work on bisimulation invariance in modal logic.
FO ∩ ADD = CFO for formulas

Lemma

Every additive FO formula $\varphi$ is $2^{\text{qr}(\varphi)}$-local

→ we can relativize every quantifier in $\varphi$ to $B(\text{free}(\varphi), 2^{\text{qr}(\varphi)})$

Lemma

Let $\varphi(x_1, \ldots, x_n)$ be an additive formula. For every $(a_1, \ldots, a_n) \in \varphi(I)$ we have $d^I(a_i, a_j) \leq (n - 1)2^{\text{qr}(\varphi)}$.

→ $\varphi \equiv \varphi \land \delta$ where $\delta = \bigwedge_{1 \leq i, j \leq n} d(x_i, x_j) \leq (n - 1)2^{\text{qr}(\varphi)}$

→ Can easily be converted to a CFO formula by pushing in $\delta$
Proposition

Every additive FO sentence $\varphi$ can be rewritten as a finite disjunction of simple local sentences

A sentence of the form $\exists x \psi(x)$ where $\psi$ is local, is called simple local

Proof idea: Rewrite $\varphi$ as $\exists x \varphi^*(x)$ where $\varphi^*$ is $(x = x) \land \varphi$

- $\varphi^*$ is invariant under disjoint copies ($\bar{a} \in \varphi^*(I)$ iff $\bar{a} \in \varphi^*(I \cup \text{cop}(I))$)
  
  $\Rightarrow \varphi$ is equivalent to a boolean combination (in DNF) of simple local sentences (Otto’s work on bisimulation invariance in modal logic)

- Eliminate negations

- Eliminate conjunctions
Example: eliminating negations

Consider an additive query $\sigma_1 \land \neg \sigma_2$ where $\sigma_1$ and $\sigma_2$ are additive.

There exists domain-disjoint instances $I$ and $J$ such that:

- $I \models \sigma_1 \land \neg \sigma_2$
- $J \models \sigma_2$

\[ I \cup J \models \sigma_1 \land \neg \sigma_2 \]

\[ J \models \sigma_2 \rightarrow I \cup J \models \sigma_2 \]
Example: eliminating conjunctions

Consider an additive query $\sigma_1 \land \sigma_2$ where $\sigma_1$ and $\sigma_2$ are additive.

There exists domain-disjoint instances $I$ and $J$

- $I \models \sigma_1$ and $I \not\models \sigma_2$
- $J \models \sigma_2$ and $J \not\models \sigma_1$

$$
I \models \sigma_1 \quad \text{additivity} \quad I \cup J \models \sigma_1 \\
\text{and} \quad I \cup J \models \sigma_1 \land \sigma_2 \quad \text{additivity} \quad \text{or} \\
J \models \sigma_2 \quad \text{additivity} \quad I \cup J \models \sigma_2
$$

Additive First-Order Queries
Connected relational algebra (CRA) is a sublanguage of RA where:
- Cartesian product is not allowed
- Equijoins are allowed

Corollary

$$RA \cap ADD = CRA$$
Guarded fragment (GF)

Define CGF as CFO formulas that are guarded

**Theorem**

\[ GF \cap ADD = CGF \]

Proof idea: If \( \varphi \in GF \cap ADD \)

1. \( \varphi \) can be written as a boolean combination (in DNF) of CGF formulas
2. Negated formulas with one free variable and positive formulas within each clause are connected
3. In each clause, the free variables of the negative part are included in those of the positive part
4. Clauses cannot contain negative sentences
5. All clauses have the same free variables
Example: positive formulas are connected

Consider $\sigma_1 \land \sigma_2$ where $\sigma_1, \sigma_2$ are additive and $\text{free}(\sigma_1) \cap \text{free}(\sigma_2) = \emptyset$

There exists domain-disjoint instances $I$ and $J$:

- Exists valuation $\alpha$ on $I$: $I, \alpha \models \sigma_1$
- Exists valuation $\beta$ on $J$: $J, \beta \models \sigma_2$
- $I \cup J, \xi \not\models \sigma_1 \land \sigma_2$ where $\xi$ equals to $\alpha$ on $\text{free}(\sigma_1)$ and to $\beta$ on $\text{free}(\sigma_2)$

\[
\begin{align*}
I, \alpha \models \sigma_1 & \xrightarrow{\text{additivity}} I \cup J, \alpha \models \sigma_1 \quad \text{and} \quad I \cup J, \xi \models \sigma_1 \\
I \cup J, \xi \models \sigma_1 \land \sigma_2 & \xrightarrow{\text{contradiction}} I \cup J, \xi \models \sigma_1 \\
J, \beta \models \sigma_2 & \xrightarrow{\text{additivity}} I \cup J, \beta \models \sigma_2 \quad \text{and} \quad I \cup J, \xi \models \sigma_2
\end{align*}
\]
Example: free variables of negative part → positive part

Consider $\sigma_1 \land \neg \sigma_2$ where $\sigma_1, \sigma_2$ are additive and that there exists $x$ in free($\sigma_2$) that is not in free($\sigma_1$)

There exists domain-disjoint instances $I$ and $J$:
- exists valuation $\alpha$ on $I$: $I, \alpha \models \sigma_1 \land \neg \sigma_2$
- $J, \beta$ is a domain-disjoint copy of $I, \alpha$

Define valuation $\xi$ that is equal to $\beta$ on $x$ and equal to $\alpha$ otherwise

\[
I, \alpha \models \sigma_1 \quad \rightarrow \quad I, \xi \models \sigma_1 \quad \xrightarrow{\text{additivity}} \quad I \cup J, \xi \models \sigma_1
\]

\[
\xi \text{ mixed for free}(\sigma_2) \quad \xrightarrow{\text{additivity}} \quad I \cup J, \xi \models \neg \sigma_2 \quad \text{and}
\]

\[
\xi \text{ mixed for free}(\sigma_1 \land \neg \sigma_2) \quad \xleftarrow{\text{contradiction}} \quad I \cup J, \xi \models \sigma_1 \land \neg \sigma_2
\]
Complexity of additivity in GF

Theorem

Additivity of guarded formulas is \(2\text{ExpTime}\)-complete

Hardness: reduction from satisfiability for guarded formulas, which is \(2\text{ExpTime}\)-hard:

\[ \varphi \text{ is unsatisfiable } \iff \varphi \land \exists x S(x) \land \exists y T(y) \text{ is additive} \]

Membership: polynomial reduction to satisfiability for guarded formulas

- Known that satisfiability is decidable in \(2^{O(n)} \cdot 2^a \log^a\) time for formulas of size \(n\) and maximum arity \(a\)
- Our reduction preserves arity
Reduction to satisfiability

Add two unary relation names $U_1$ and $U_2$ to the schema $S$

- Use $U_1$ and $U_2$ to mimic domain-disjoint instances

Consider instances $I$ where the $U_1$ - and $U_2$-facts partition $I$ into domain-disjoint instances $(I_1, I_2)$

- Definable by a GF formula $\varphi_{cons}$

For a GF formula $\varphi$ define GF formulas $\psi_i$ for $i = 1, 2$:

- Relativize the free variables and quantifiers in $\varphi$ to $U_i$

$$\varphi(l_1) \cup \varphi(l_2)$$

Lemma

$$\varphi \text{ is additive } \iff \varphi_{cons} \land \neg(\varphi \leftrightarrow (\psi_1 \lor \psi_2)) \text{ is unsatisfiable}$$

Additive First-Order Queries
Positive Existential (PE)

Fragment of FO where only $\exists$, $\land$ and $\lor$ are allowed

Theorem (earlier work)

\[ PE \cap \text{ADD} = \text{CUCQ} \]

Deciding additivity for UCQs is NP-complete (earlier work)

Translation from PE to UCQ takes exponential time

- Gives an exponential upper bound for additivity checking

We have shown a more precise upper bound

Theorem

Additivity of PE formulas is $\Pi^P_2$-complete. The lower bound holds even over unary and binary relations.
Future work

- Investigate the complexity of the shortest connected formula equivalent to an additive formula
- Our hardness proof for additivity in GF is not specific to GF. Interesting to apply the reduction to other decidable fragments
- Our version for additivity is very strict in the context of distributed computation
  - Interesting to consider other ways to partition instances