



Polynomial bounds on clique-width in terms of tree-width

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Reference : B.C.: From tree-decompositions to clique-width terms, *Discrete Applied Maths*, 2017, in press (on line on ScienceDirect.com).

Various articles cited below.

Motivation

Fixed-parameter tractable (FPT) graph algorithms for monadic second-order (MSO) expressible problems,
for graphs of bounded **tree-width** (**twd**) or **clique-width** (**cwd**),
based on **automata** running on algebraic terms denoting the
(decomposed) input graphs.

Tools: **Fly-automata** (FA): they *compute* their transitions, to
overcome the “huge size problem”,

Tree-decompositions encoded by clique-width terms,

Linear bounds on **cwd** in terms of **twd** for sparse graphs.

The basic theorem : Each MSO property of graphs of **cwd** or **twd** at most **k** is decidable in time **$f(k)$** x ***number of vertices***.

Facts: Extends to MSO properties expressed with **edge set quantifications**, for graphs of bounded tree-width (***not*** bounded **cwd**).

Graphs given with relevant decompositions, of “small width”.

Optimal decompositions are difficult to construct (NP-complete problems). But optimality is not essential.

Computation of graph evaluations

$P(\underline{X})$ is a property of tuples \underline{X} of sets of vertices (usually MSO expressible).

$\exists \underline{X}.P(\underline{X})$: the basic, “Boolean evaluation”.

$\# \underline{X}.P(\underline{X})$: number of satisfying tuples \underline{X} .

Sp $\underline{X}.P(\underline{X})$: **spectrum** = the set of tuples of cardinalities of the components of the tuples \underline{X} that satisfy $P(\underline{X})$.

MinCard $\underline{X}.P(\underline{X})$: minimum cardinality of \underline{X} satisfying $P(\underline{X})$.

Computations using fly-automata (by Irène Durand)

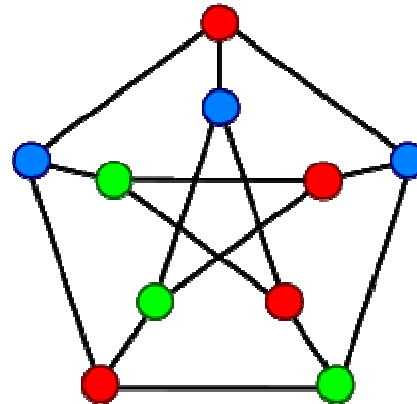
Number of 3-colorings of the 6×525 rectangular grid (of clique-width 8) in 10 minutes.

4-acyclic-colorability of the **Petersen graph** (clique-width 5) in 1.5 minutes.

(3-colorable but not acyclically;

red and **green** vertices

induce a cycle).



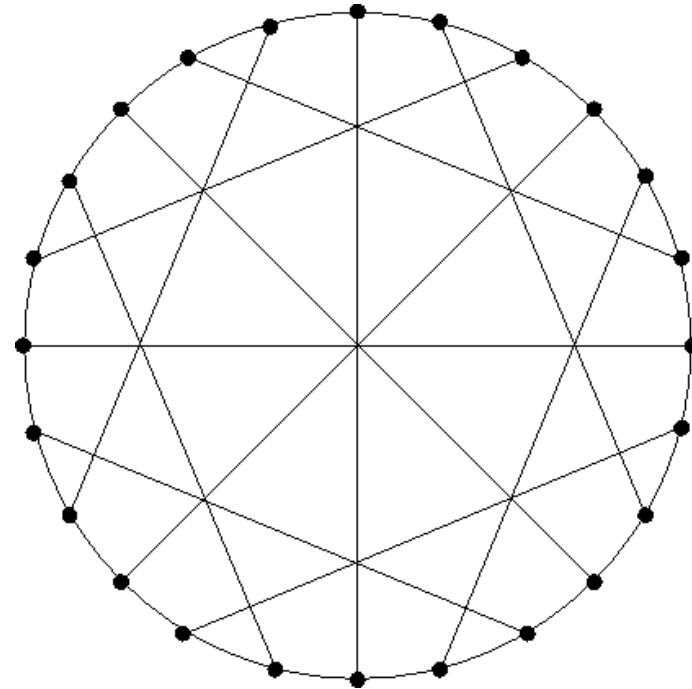
The **McGee graph**

is defined by a clique-width term
of size 99 and depth 76.

This graph is 3-acyclically colorable.

Checked in 40 minutes.

Even in 2 seconds by enumerating the accepting
runs, and stopping as soon as a successful one is found.



Informal review of definitions and basic facts.

1) **Graphs** are finite, simple, loop-free, directed or not.

A graph G can be given by the logical structure

$$(V_G, \text{edg}_G(\cdot, \cdot)) = (\text{vertices, adjacency relation})$$

2) Alternative description of graphs :

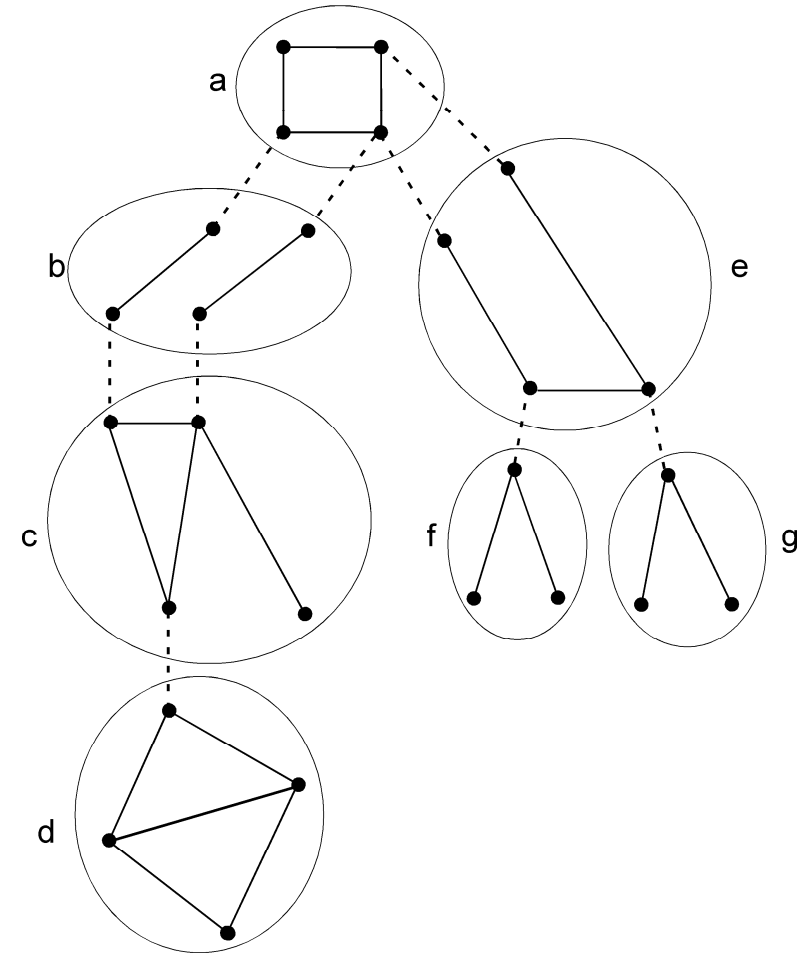
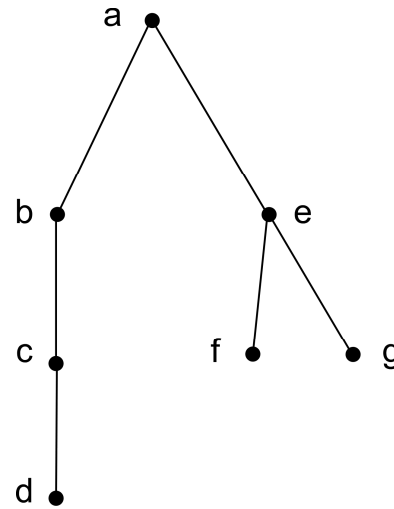
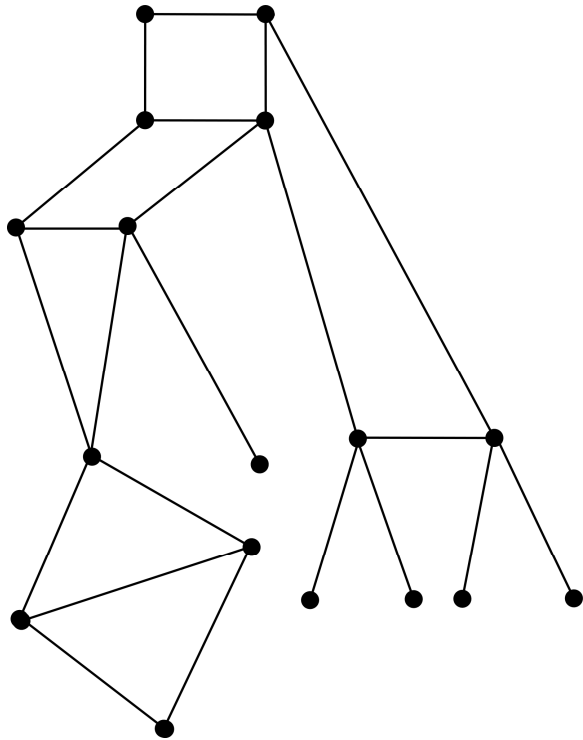
$$\text{Inc}(G) := (V_G \cup E_G, \text{inc}_G(\cdot, \cdot))$$

$$= (\text{vertices *and edges*, incidence relation})$$

→ the bipartite *incidence graph* of G .

MSO formulas on $\text{Inc}(G)$ can use quantifications on sets of edges of the considered graph G .

3) Tree-width ($\text{twd}(G)$) is well-known.



width of decomposition : 3
 dotted lines : equal vertices

4) Clique-width : algebraic construction of graphs

Vertices are labelled by a, b, c, \dots . A vertex labelled by a is an a -vertex.

Binary operation: disjoint union : \oplus

Unary operations: edge addition denoted by $Add_{a,b}$

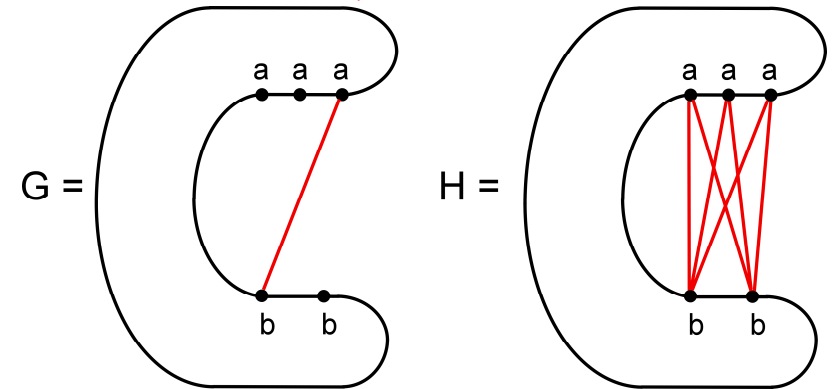
$Add_{a,b}(G)$ is G augmented

with (un)directed edges from (between) every a -vertex to (and) every b -vertex.

vertex relabellings :

$Relab_a \rightarrow b(G)$ is G with every a -vertex is made into a b -vertex

Basic graphs : a denotes a vertex labelled by a



The **clique-width** of G , denoted by $\text{cwd}(G)$, is the smallest k such that G is defined by a term using k labels.

Such a term is a decomposition of G as a gluing of complete bipartite graphs. k indicates the “complexity of gluings”, not size of components.

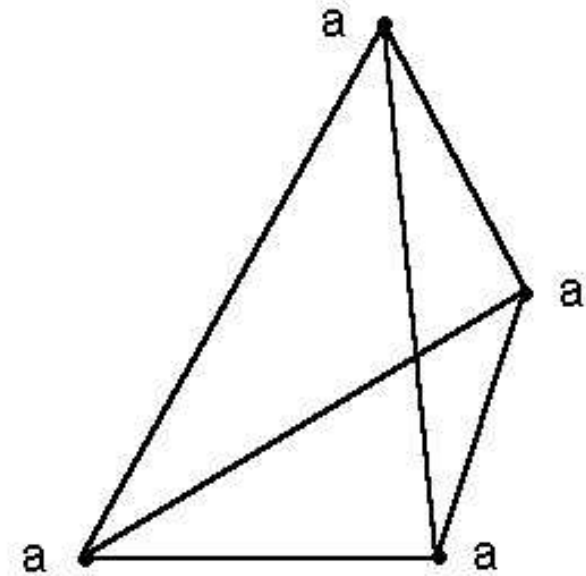
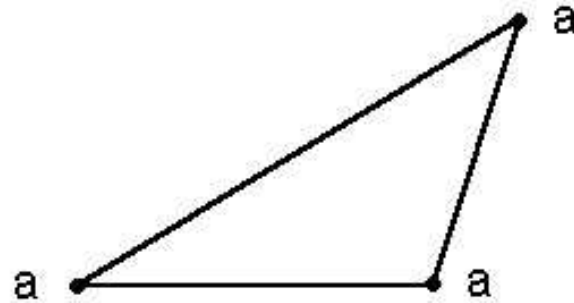
Classes of bounded clique-width:

cographs, cliques, complete bipartite graphs, trees,
any class of bounded tree-width.

Classes of unbounded clique-width:

Planar graphs, chordal graphs.

Example 1 : Cliques (**a**-labelled) have clique-width 2 and unbounded tree-width.



K_n is defined by t_n where $t_1 = \mathbf{a}$

$$t_{n+1} = \text{Relab } b \rightarrow a (\text{Add } a, b (t_n \oplus \mathbf{b}))$$

Example 2 : Cographs (**a**-labelled) are generated by \oplus and \otimes defined

$$\begin{aligned} \text{by: } G \otimes H &= \text{Relab } b \rightarrow a (\text{Add } a, b (G \oplus \text{Relab } a \rightarrow b (H))) \\ &= G \oplus H \text{ with "all edges" between } G \text{ and } H. \end{aligned}$$

Remark : An algebraic expression of tree-width is possible, by using *parallel composition* $G // H$ instead of disjoint union $G \oplus H$.

This operation glues G and H by fusing, for each label a , the (*unique*) a -vertex of G and the (*unique*) a -vertex of H .

But the construction of an automaton running on terms over $//$ denoting graphs G of $\text{twd} \leq k$ intended to check an MSO property of $\text{Inc}(G)$ is more complicated because of these fusions. The basic fact for \oplus is : $G \oplus H \models \varphi(X)$ if and only if

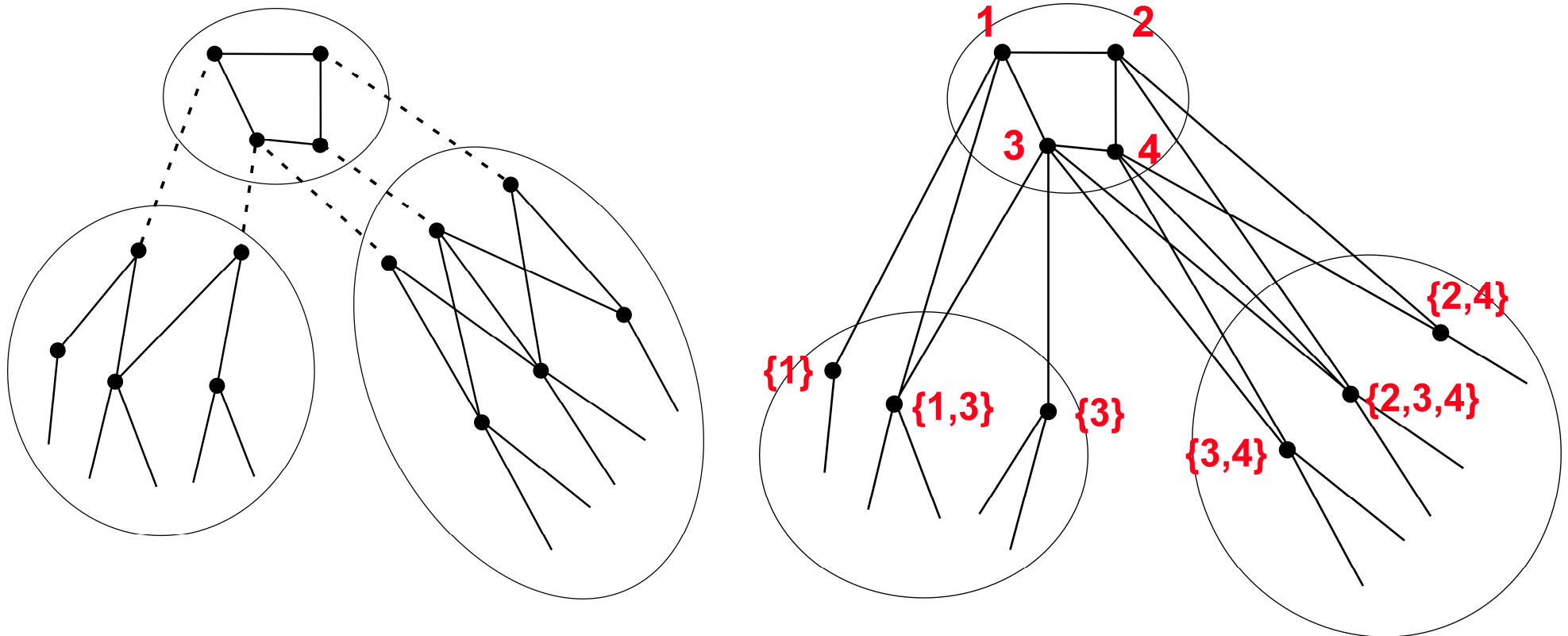
$$G \models \psi_1(X \cap V_G) \text{ and } H \models \theta_1(X \cap V_H)$$

$$\text{or } G \models \psi_2(X \cap V_G) \text{ and } H \models \theta_2(X \cap V_H) \dots$$

$$\text{or } G \models \psi_p(X \cap V_G) \text{ and } H \models \theta_p(X \cap V_H)$$

Comparing tree-width and clique-width (undirected graphs)

$\text{cwd}(G) \leq 3 \cdot 2^{\text{tw}(G) - 1}$ (Corneil & Rotics, the exponential is **not** avoidable)



If a box of the tree-dec has k vertices, then $2^k - 1$ labels may be necessary to specify how the vertices below it are linked to its vertices.

We have : $\text{cwd}(G) \leq 3 \cdot 2^{\text{tw}(G) - 1}$

Tree-width is not bounded from clique-width (consider cliques).

However : $\text{tw}(G) < 3(r-1) \cdot \text{cwd}(G)$ if G has no subgraph $K_{r,r}$

We will only consider classes of graphs with no $K_{r,r}$:

From them : **bounded tree-width \Leftrightarrow bounded clique-width.**

Question : For which classes do we have $\text{cwd}(G) = O(\text{tw}(G)^c)$ for fixed c , with “good values” of c and of hidden constants ?

Graph class	$\text{cwd}(G)$ where $k = \text{twd}(G)$
No subgraph $K_{r,r}$	$4 / (r-1)! \cdot k^r$ where $k > 2r$
Uniformly q -sparse : at most $q \cdot n$ edges for n vertices (or q -degenerated)	$2 / (q-1)! \cdot k^q$ where $k > 2q$
Nowhere dense	$f(\varepsilon) \cdot k^{1+\varepsilon}$ for each $\varepsilon > 0$
Bounded expansion or even $\diamond_1(G) \leq b$ ("nabla" ₁)	$4^b \cdot k$
no $K_{r,r}$ minor	$a^{r \log \log r} \cdot k$
p -planar	$h(p) \cdot k$

Large functions f, g, h and value a !

With “good” bounds :

1-planar	$18k - 29$ (undirected)
planar	$6k - 2$ ($32k - 24$ if directed)
degree $< d$	$k.d + 1$ (also directed)
incidence graph	$k + 3$ ($2k + 4$ if directed)
grids	$k + 2$

Remark: About incidence graphs of graphs of bounded **tree-width** and MSO_2 properties.

MSO_2 means expressed by an MSO formula using edge set quantifications.

Example: There exists a set of edges forming a perfect matching, or forming a Hamiltonian path. Not possible without such quantifications.

- 1) From of a tree-decomposition of G of width k , we construct a clique-width term t for $\text{Inc}(G)$ of “small” width $k+3$ (or $2k+4$); **no exp. !**
- 2) We translate an MSO_2 formula φ for G into an MSO formula θ for $\text{Inc}(G)$.
- 3) The corresponding automaton $A(\theta)$ takes term t as input.

Proof method for making tree-decompositions into **cwd** terms

For a graph G and Y a set of vertices :

$\mu_G(Y) :=$ the number of sets $N_G(x) \cap Y$ for $x \notin Y$. ($N_G(x)$: neighbours of x)

More generally, **neighbourhood complexity** :

$\nu_G^r(Y) :=$ the number of sets $N_G^r(x) \cap Y$.

($N_G^r(x)$: neighbours at distance at most r of x)

Hence : $\mu_G(Y) \leq \nu_G^1(Y)$.

Also : $\mu_G(Y) \leq (4^r + 4r) \cdot |Y|$ where $r = \diamond_1(G)$ (Gajarsky et al.)

Lemma : If $\text{twd}(G) \leq k$, and $\mu_G(Y) \leq m$ whenever $|Y| \leq k + 1$,
then $\text{cwd}(G) \leq m + 1$.

For each graph class, we bound $\mu_G(Y)$ in terms of $|Y|$.

For planar graphs, we use the bound $3n - 6$ on the number of edges ;

for q -sparse graphs, we use an orientation of indegree at most q .

In all cases we transform a tree-decomposition into a clique-width term based on the same tree.

Proof sketch for planar graphs.

Enough to consider a bipartite graph with vertex set $X \cup Y$ and $|Y| = k$.

There are at most $k+1$ sets $N_G(x) \cap Y$ for x of degree 0 or 1, ($x \in X$).

There are at most $3k-6$ sets $N_G(x) \cap Y$ for x of degree 2 : each of them corresponds to an edge of a planar graph with vertex set Y .

There are at most $2k-4$ vertices x of degree > 2 : let Z be these vertices : $3 \cdot |Z| \leq |E| \leq 2 \cdot (|Z| + k) - 4$ (planar bipartite).

$$\text{Total : } k+1 + 3k-6 + 2k-4 = 6k - 9.$$

Graph classes of bounded expansion

(Nesetril, Ossona de Mendez)

Some cases : bounded degree, s -colorable,
minor closed, hence planar or bounded tree-width,
topologically closed (by contracting paths), p -planar.

Definition : A class \mathcal{C} has bounded expansion if

$\forall d \exists c \forall G \in \mathcal{C}$ and H a d -shallow minor of G ,

we have $|E_H| \leq c \cdot |V_H|$

d -shallow minor : contracting connected subgraphs of radius $\leq d$.

Notation:

$$\diamond_d(G) := \text{Max} \{ |E_H| / |V_H|, H \text{ } d\text{-shallow minor} \} \leq c$$

Theorem (Reidl et al.) :

A class C has bounded expansion \Leftrightarrow for each r , we have :

$$\exists c \forall G \in C \forall Y \subseteq V_G : v_G^r(Y) \leq c \cdot |Y|$$

Hence, if C has bounded expansion (take $r = 1$) :

$$\text{cwd}(G) = O(\text{twd}(G)) \text{ for all } G \in C$$

Whence, the answer for p -planar graphs (B.E. by Nešetřil et al.)

But the bound c for $r = 1$ is huge.

Remark 1 : Several bounds for p -planar graphs:

$\text{cwd}(G) \leq h(p) \cdot \text{twd}(G)$ with huge $h(p)$.

$\text{cwd}(G) \leq (6 \cdot \text{twd}(G))^{p+1}$ because a p -planar graph is the union of $\leq p+1$ planar graphs.

$\text{cwd}(G) \leq (18 \cdot \text{twd}(G))^{p/2+1}$ because a p -planar is the union of $\leq p/2+1$ 1-planar graphs.

$\text{cwd}(G) \leq a \cdot \text{twd}(G)^q$ where $q = 4 \cdot p^{1/2}$

because p -planar graphs are uniformly q -sparse (Pach and Toth).

Remark 2 :

For a graph of tree-width d , given by a non-optimal tree-decomposition of width k , we obtain a clique-width term of width at most $f(d).k$ for some fixed function f .

The results of Corneil and Rotics give $f(d) \geq 2^{d/2} / d$.

To be done : “good” estimation of $f(d)$.

More on sparse graphs : **nowhere dense** graph classes

They include : bounded expansion, *locally* bounded expansion
and thus locally bounded tree-width.

Definition (Nesetril, Ossona de Mendez) : A class C is **nowhere dense** if

$\forall d \exists c \forall G \in C \forall H$ **d -shallow minor** of G :

$\omega(H) \leq c$ ($\omega(H)$:= max size of a clique in H).

Theorem (Eickmeyer et al.) : A class C is *nowhere dense* \iff

for each r and $\varepsilon > 0$:

$$\exists c \forall G \in C \forall Y \subseteq V_G : v_G^r(Y) \leq c \cdot |Y|^{1+\varepsilon}$$

Hence, if C is nowhere dense (take $r = 1$), then for each $\varepsilon > 0$:

$$\text{cwd}(G) = O(\text{twd}(G)^{1+\varepsilon}) \text{ for all } G \in C$$

Proper inclusions of classes :

1. Graphs without $K_{r,r}$ as a subgraph are **not** uniformly q -sparse for any q (Erdős & Stone, see Diestel's book), but $\text{twd}(G) = O(\text{cwd}(G))$ (Gurski & Wanke).

2. The class of graphs G that are uniformly q -sparse, *i.e.* each subgraph H with n vertices has $\leq q \cdot n$ edges, is **not** nowhere dense : consider cliques with subdivided edges.

We only have, presently, $\text{cwd}(G) = O(\text{twd}(G)^q)$.

Question : Do we have $\text{cwd}(G) = O(\text{twd}(G)^{q/2})$?

3. The class of graphs whose degree is at most the girth is nowhere dense by **does not have** bounded expansion. (N. & OdM's book)

Discussion

- in many cases the constants are “bad” (exponential) ;
- however, they are **not reached**, or in weird cases only ;
- better bounds (cf. the cases of planar and 1-planar graphs)
should be determined for classes of particular interest ;
- the algorithm given below works for arbitrary
tree-decompositions (given **by normal trees**) as input ;
the time and space need not be huge (no “search”).

Some open questions

“Good” bounds for p -planar graphs and other classes of bounded expansion.

Where lies the class of q -quasi planar graphs (have drawings where no set of q edges that cross pairwise) ?

It is known that q -planar graphs are $(q+1)$ -quasi planar and that q -quasi planar graphs have $n \cdot (\log n)^{O(\log q)}$ edges for n vertices. For q at most 4, they are uniformly r -sparse.

An algorithm to transform tree-decompositions
into clique-width terms.

First : How to specify tree-decompositions ?

Instead of the classical definition (T, f) , we use *partial k-trees* in the following way. A *normal tree* T for a graph G is :

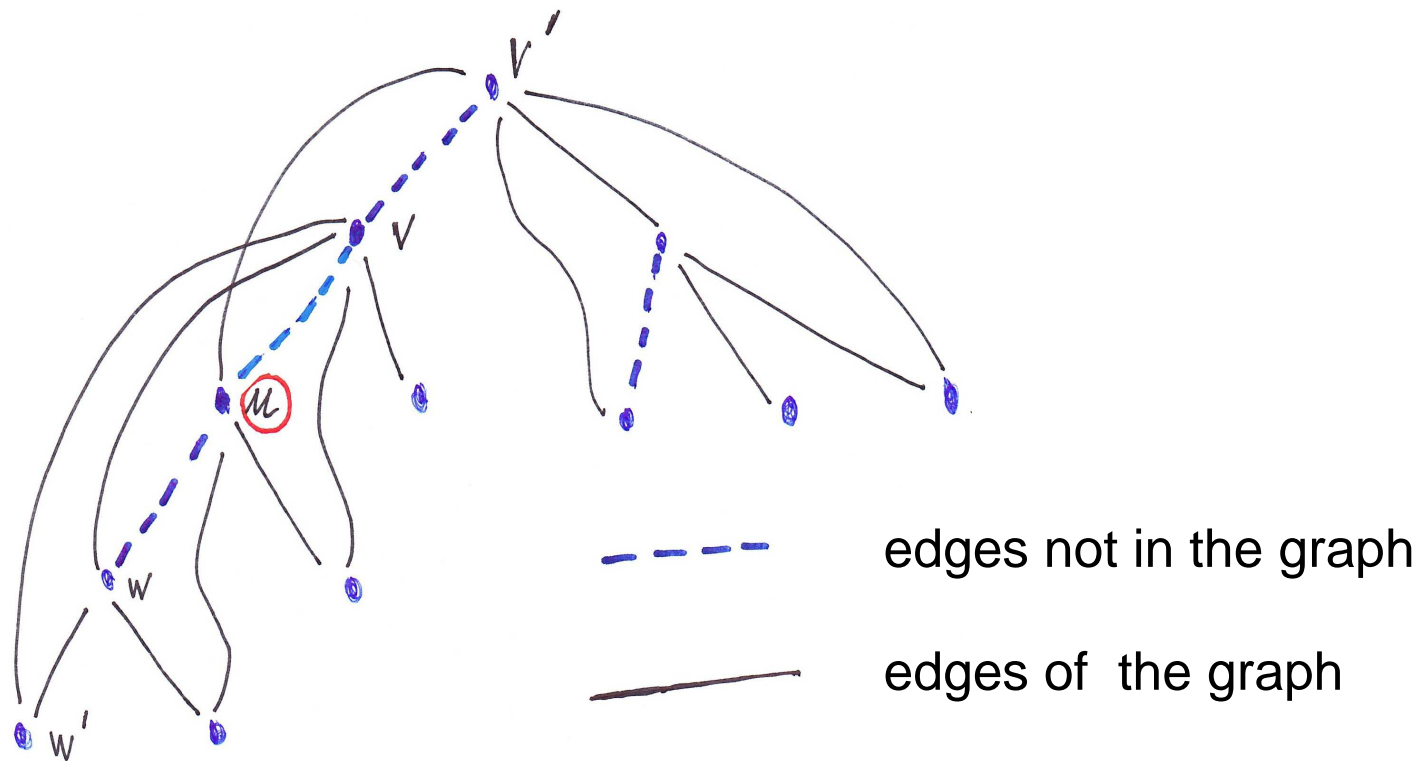
rooted, its nodes *are the vertices* of the graph and adjacent vertices of the graph are *comparable* for the ancestor relation of T .

Then T is the tree of a tree-decomposition (T, f_T) where :

$$f_T(u) := \{u\} \cup \{v >_T u \mid v \text{ is adjacent to some } w \leq_T u\}.$$

Every tree-dec (T', f) can be made (T, f_T) of same width for a normal tree T (by contracting edges in T' and inserting nodes on edges; no complicated transformation).

$f_T(u) := \{u, v, v'\}$: the edges from w, w' “jump” over u



Remarks :

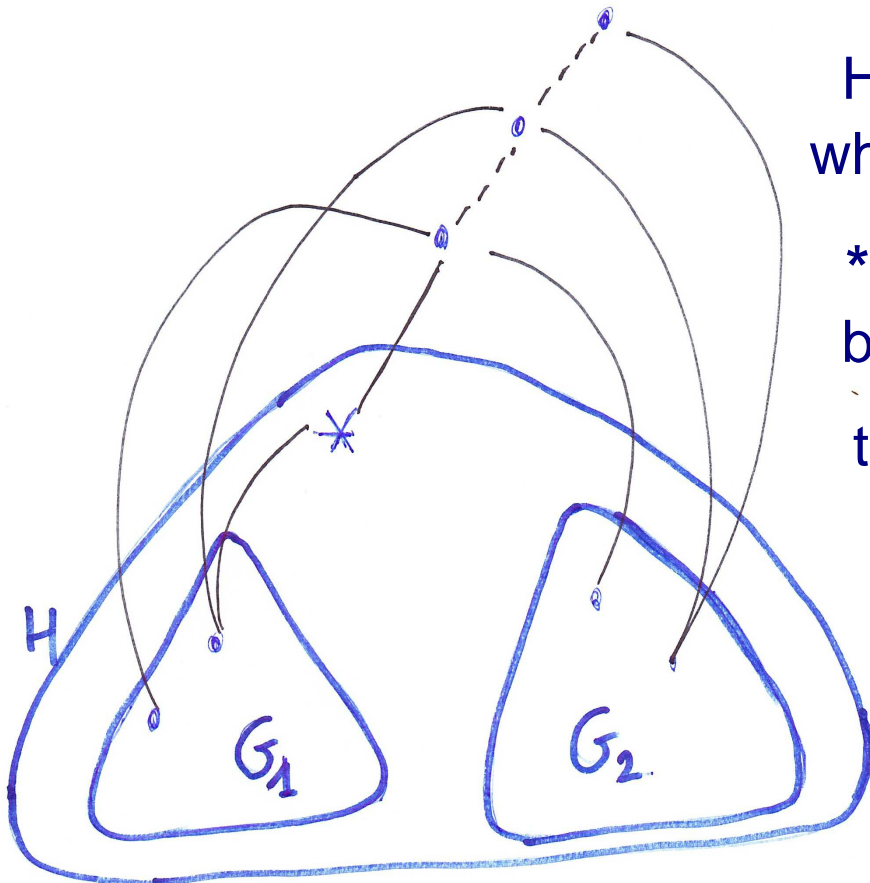
1. We get a *compact data structure* for the **graph** and a **tree-decomposition** : $R = (V_G, \text{edg}_G, \text{parent}_T)$

from which f_T (describing the “bags”) is easily computable.

2. This triple is also a *convenient logical structure* : the bags can be described by an MSO formula $\varphi(u, X)$ saying “ $X = f_T(u)$ ” (in R)

3. This description corresponds to the notion of a **partial k-tree**, obtained by edge deletions from a **k-tree**.

Bottom-up inductive construction of a clique-width term from a normal tree-decomposition.



$H = \text{RELAB} (\text{ADD}(G_1 \oplus G_2 \oplus *))$
 where ADD adds the edges between
 $*$ and the vertices in $G_1 \oplus G_2$, on the
 basis of the labels in $G_1 \oplus G_2$
 that encode subsets of $f_T(*)$.
 The **number of such labels** is
 bounded by $\mu_G(f_T(*))$.
 RELAB : relabellings to update the
 labels in $G_1 \oplus G_2$ and change $*$ into
 the correct label for H .

Remarks :

1. The algorithm uses time $O(n \cdot k \cdot m^2)$ where :

n = number of vertices (nodes of the tree T),

k = the width of the tree-decomposition (T, f_T) ,

m = number of labels of the produced clique-width term.

2. In this construction, $\text{add}_{a,b}$ only creates “stars” $K_{1,p}$, but no $K_{q,p}$. The full power of edge addition is not used. We do not get optimal clique-width terms (as examples can show).

Conclusion:

From a “good” tree-decomposition of a sparse graph (planar, bounded degree, etc...), we can get, in some cases, a “good” clique-width term, of comparable width (avoiding the general exponential jump).

There are many algorithms that construct “good” (not optimal) tree-decompositions, but not so many that construct “good” clique-width terms. Algorithms based on rank-width do not give “good terms”.

Clique-width terms yield easier constructions of fly-automata than tree-decompositions.

Fly-automata for the verification of MSO graph properties

Standard proof of the basic theorem: one constructs, for each MSO formula φ and integer k , a finite automaton $A(\varphi, k)$ that takes as input a term denoting a graph G of clique-width $\leq k$ and answers in time $f(k) \cdot n$ whether $G \models \varphi$ (where n is the number of vertices).

The construction is by induction on the structure of φ .

Difficulty : The *finite* automaton $A(\varphi, k)$ is too large to be implemented by a transition table as usual as soon as $k \geq 2$: $2^{(2^{(\dots 2^k \dots)})}$ states, because of quantifier alternations.

To overcome this difficulty, we use *fly-automata* whose states and transitions are *described* and *not tabulated*. Only the (say 100) transitions necessary for an input term (say of size 100) are computed “on the fly”.

Sets of states can be infinite and fly-automata can compute values, for example, the number of *p-colorings* of a graph.

Fly-automaton (FA)

Definition : $A = \langle F, Q, \delta, \text{Out} \rangle$ (FA that computes a function).

F : finite **or countable** (**effective**) set of operations,

Q : finite **or countable** (**effective**) set of states (integers, pairs of integers, *etc.* : states are encoded by finite words),

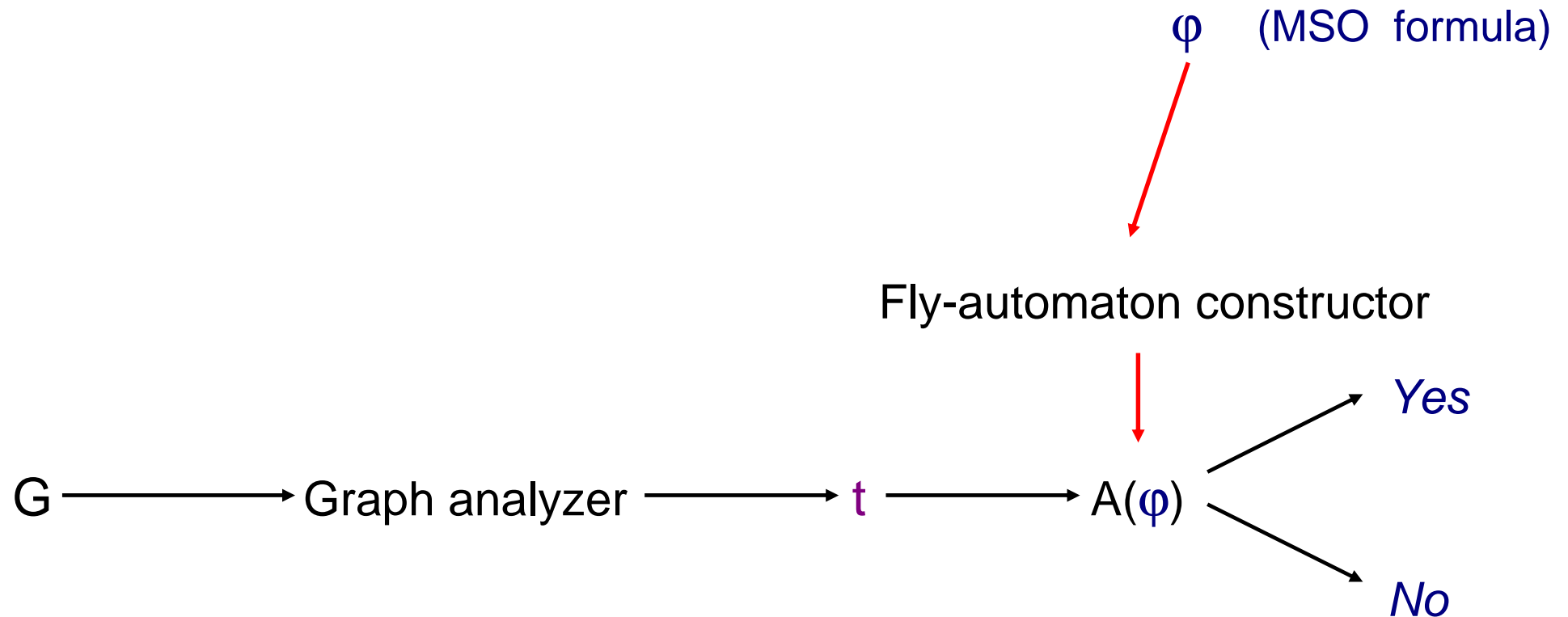
Out : $Q \rightarrow D$, **computable** (D is an effective set, coded by finite words).

δ : **computable** (bottom-up) transition function

Nondeterministic case : δ is **finitely multi-valued**. **Determinization** works.

An FA defines a **computable function** : $T(F) \rightarrow D$, hence, a **decidable property** if $D = \{True, False\}$.

The MSO meta-theorem through *fly-automata*



$A(\varphi)$: *a single infinite fly-automaton*. The time taken by $A(\varphi)$ is $f(k).n$ where k depends on the operations occurring in t and bounds the tree-width or clique-width of G .

Computation time of a fly-automaton (FA)

F : all clique-width operations, F_k : those using k labels.

On term $t \in T(F_k)$ defining $G(t)$ with n vertices, if a fly-automaton takes time bounded by :

$(k + n)^c \rightarrow$ it is a P-FA (a polynomial-time FA),

$f(k) \cdot n^c \rightarrow$ it is an FPT-FA,

a. $n^{g(k)} \rightarrow$ it is an XP-FA.

The associated algorithm is polynomial-time, FPT or XP for clique-width as parameter. (The important notion is the max. size of a state.)

All dynamic programming algorithms based on clique-width terms can be described by FA.

Fly-automata can be constructed :

- either “directly”, from our understanding of the considered graph properties,
- or “automatically” from a logical description,
- or by combining previously constructed automata.

Example of a direct construction for p -coloring :

Checking that a “guessed” p -coloring is good: a state is a set of pairs (a, j) where a is a label and j is a color (among $1, \dots, p$) or Error.

Checking the existence of a good p -coloring : a set of such states, in practice not of maximal (exponential) size.

Combinations and transformations of fly-automata.

Product of A and B : states are pairs of a state of A and one of B.

Determinization of A : states of $\text{Det}(A)$ are finite sets of states of A because the transition is *finitely* multi-valued. *At each position in the term, $\text{Det}(A)$ gives the finitely many states that can in some computation (the automaton A can be infinite).*

Counting determinization of A, yielding $\text{CDet}(A)$:

a state of $\text{CDet}(A)$ is a finite multi-set of states of A (giving the *number of runs* that can yield a state of A, not only the existence).

Inductive construction for $\exists \underline{X}. \varphi(\underline{X})$ with $\varphi(\underline{X})$ MSO formula

Atomic formulas (for example $X \subseteq Y$, $\text{edg}(X, Y)$) : direct constructions

$\neg P$ (negation) : as FA are run deterministically (by computing at each position the **finite** set of reachable states), it suffices to exchange accepting and non-accepting states.

$P \wedge Q$, $P \vee Q$: products of automata.

How to handle free variables for **queries** $\varphi(\underline{X})$ and for $\exists \underline{X}. \varphi(\underline{X})$?

Terms are equipped with **Booleans** that encode assignments of vertex sets V_1, \dots, V_p to the free set variables X_1, \dots, X_p of MSO formulas (*formulas are written without first-order variables*):

1) we replace in F each **a** by the nullary symbol

(a, (w₁, ..., w_p)), $w_i \in \{0, 1\}$: we get $F^{(p)}$ (*only nullary symbols are modified*);

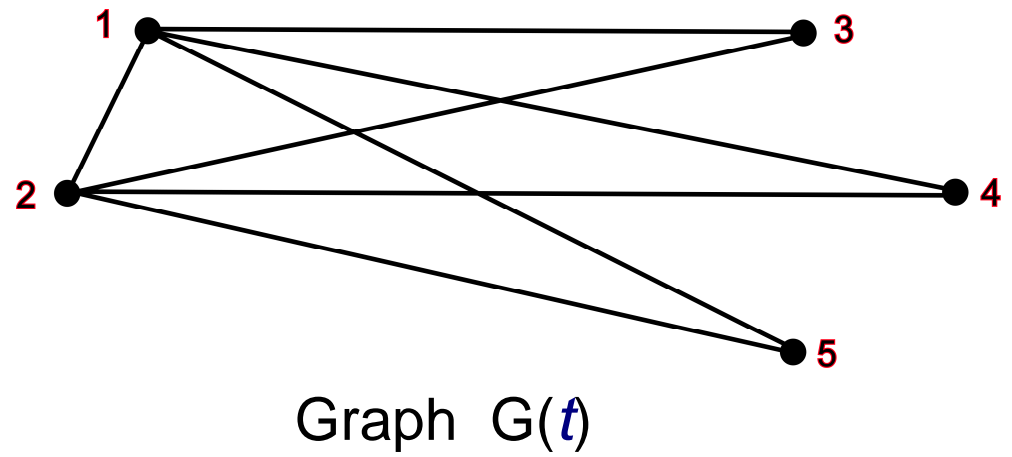
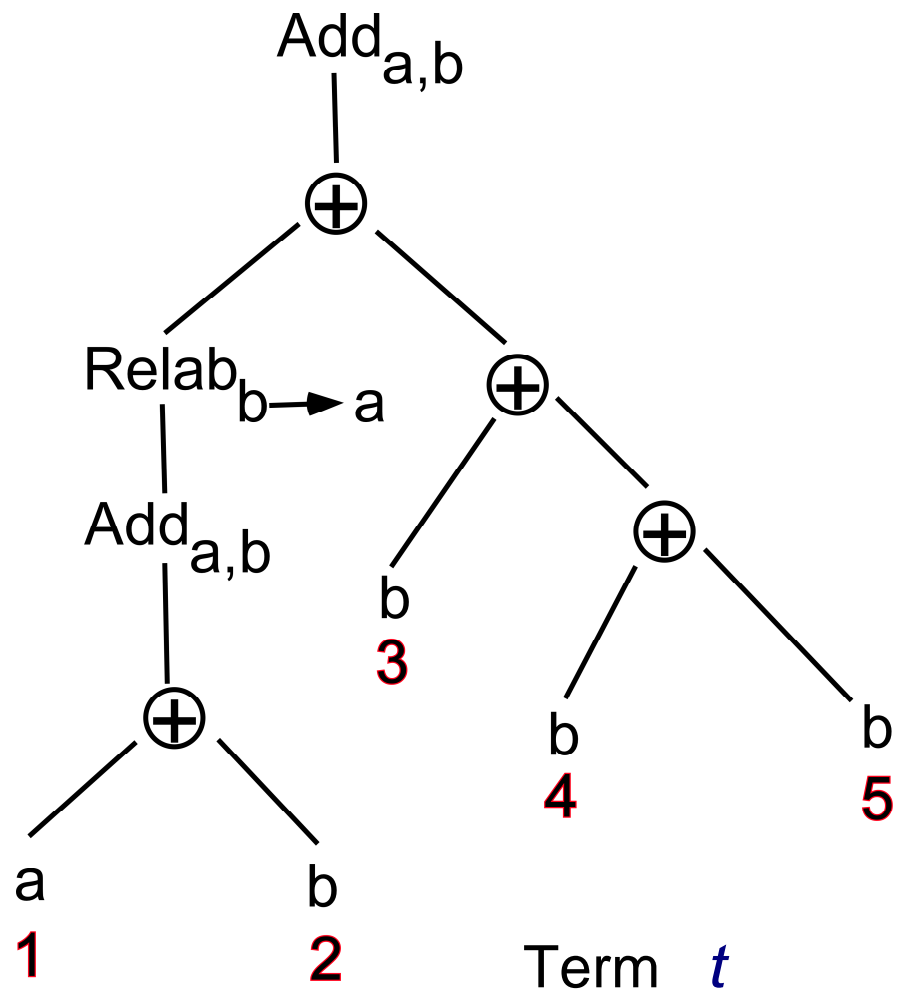
2) a term **s** in $\mathbf{T}(F^{(p)})$ encodes a term **t** in $\mathbf{T}(F)$ and an assignment of sets V_1, \dots, V_p to the set variables X_1, \dots, X_p :

if **u** is an occurrence of **(a, (w₁, ..., w_p))**, then

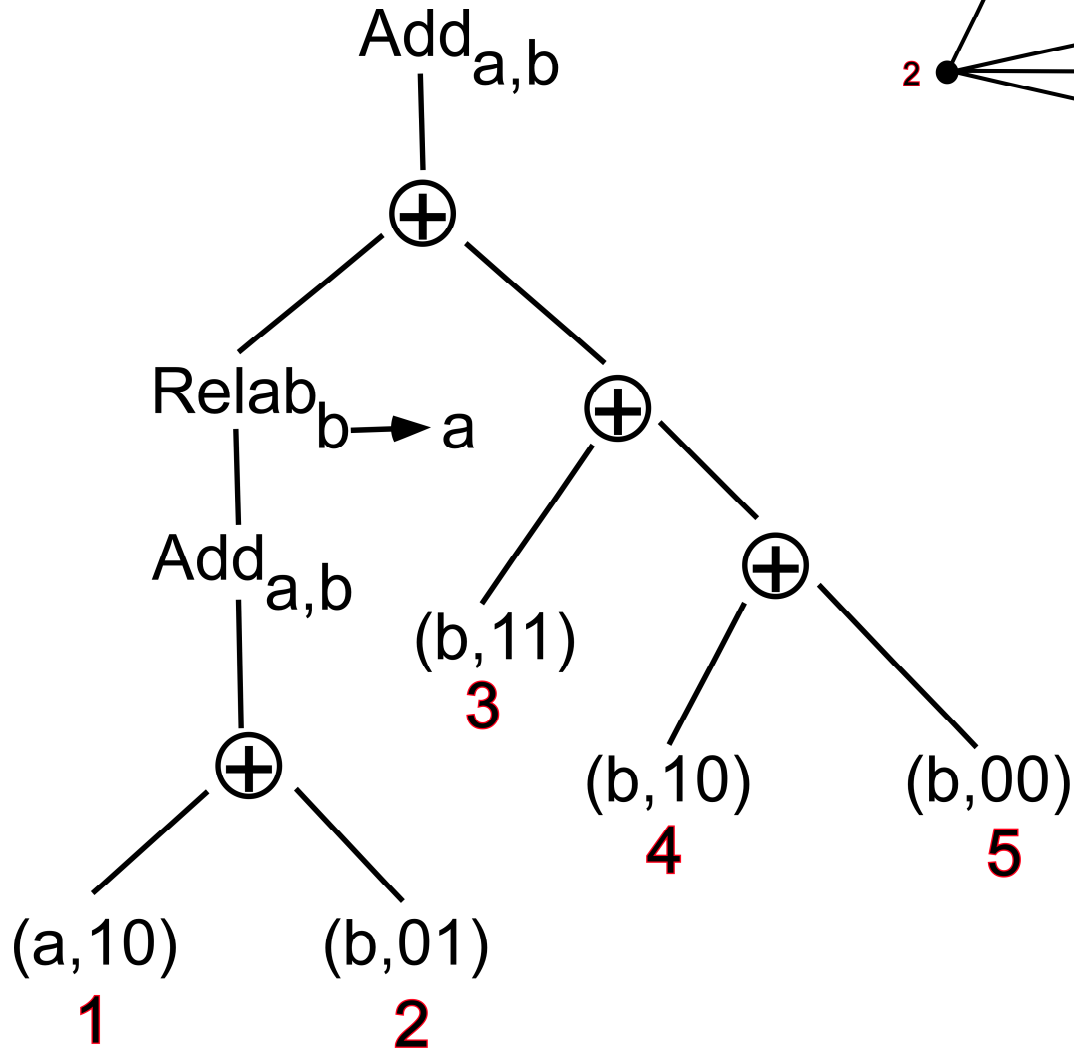
$w_i = 1$ if and only if **u** $\in V_i$.

3) **s** is denoted by **t*** (V_1, \dots, V_p)

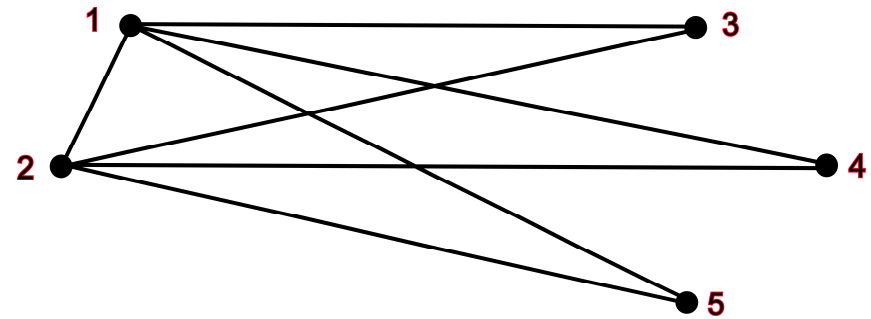
Example



Example (continued)



Term $t^*(V_1, V_2)$



$$V_1 = \{1,3,4\}, V_2 = \{2,3\}$$

By an induction on φ , we construct, for each $\varphi(\underline{X})$, $\underline{X}=(X_1,\dots,X_p)$, a fly-automaton $A(\varphi(\underline{X}))$ that recognizes:

$$L(\varphi(\underline{X})) := \{ t^* (V_1, \dots, V_p) \in \mathbf{T}(F^{(p)}) \mid (G(t), V_1, \dots, V_p) \models \varphi \}$$

Quantifications: Formulas are written without \forall

$$L(\exists X_{p+1} . \varphi(X_1, \dots, X_{p+1})) = \text{pr}_{p+1}(L(\varphi(X_1, \dots, X_{p+1})))$$

$$A(\exists X_{p+1} . \varphi(X_1, \dots, X_{p+1})) = \text{pr}_{p+1}(A(\varphi(X_1, \dots, X_{p+1})))$$

where pr_{p+1} is the *projection* that eliminates the last Boolean;

→ a *non-deterministic* FA denoted by $\text{pr}_{p+1}(A(\varphi(X_1, \dots, X_{p+1})))$, to be run deterministically.

Remark: If a graph is denoted by a clique-width term t , then each of its vertices is represented in t at a **single position** (an occurrence of a nullary symbol).

If the operation $//$ is also used ($G // H$ is obtained from disjoint G and H by fusing some vertices of G to some vertices of H , in a precise way fixed by labels), then a vertex of $G//H$ is represented by **several positions** of the term. The automaton that checks a property $\varphi(X_1, \dots, X_p)$ of G denoted by a term t must also check that the Booleans that specify (X_1, \dots, X_p) agree on all positions of t that specify a same vertex of G .

We have no such difficulty if we use disjoint union instead of $//$. Hence, for representing tree-decompositions, clique-width terms may be more convenient.

Application to MSO_2 properties of graphs of bounded tree-width *via* incidence graphs.

1) *Recall* : From of a tree-decomposition of G of width k , we construct a term t for $\text{Inc}(G)$ of “small” clique-width $k+3$ (or $2k+4$).

2) *Recall* : We translate an MSO_2 formula φ for G into an MSO formula θ for $\text{Inc}(G)$.

3) The corresponding automaton $A(\theta)$ takes term t as input. But an atomic formula $\text{edg}(X,Y)$ of φ is translated into $\exists U. \text{inc}(X,U) \wedge \text{inc}(U,Y)$ in θ which adds one level of quantification.

Fact : The automaton $A(\theta)$ remains manageable.

For certain graph properties P , for example “connectedness”, “contains a directed cycle” or “outdegree $< p$ ”, we have :

$$P(G) \Leftrightarrow P(\text{Inc}(G)).$$

The automaton for graphs G defined by clique-width terms can be used “directly” for the clique-width terms that define the graphs $\text{Inc}(G)$.

Summary : Checking properties of G of **tree-width** $< k$

MSO property	MSO ₂ property
cwd term for G of width $O(k)$ or $O(k^q)$ in “ good cases ” and exponential in bad ones	cwd term for Inc (G) of width $O(k)$; more complicated automaton in some cases, because of $\text{edg}(X,Y)$

General conclusion

- 1) By uniform constructions, we get dynamic programming algorithms based on **fly-automata**, that can be quickly constructed from logical descriptions → *flexibility*. A small modification of the input formula is reflected easily in the automaton.
- 2) It is hard to obtain upper-bounds to time computations. We do not get better algorithms than the specific ones that have been developed.

- 3) Even for graphs given by tree-decompositions, clique-width terms are appropriate because of two facts:
- (a) Fly-automata are **simpler to construct** and
 - (b) it is **practically possible** to translate tree-decompositions of “certain” sparse graphs into clique-width terms.
- 4) Fly-automata are **implemented**. Tests have been made for colorability and connectedness problems.