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Distributed Computing

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Principles of Distributed Computing
Exercise 14: Sample Solution

1 Communication Complexity of Set Disjointness

a) We obtain

$$M^{DISJ} = \left(\begin{array}{c|cccccccc c}
\text{DISJ} & 000 & 001 & 010 & 011 & 100 & 101 & 110 & 111 & \leftarrow x \\
\hline
000 & 1 & 1 & 1 & 1 & 1 & 1 & 1 & 1 & \\
001 & 1 & 0 & 1 & 0 & 1 & 0 & 1 & 0 & \\
010 & 1 & 1 & 0 & 0 & 1 & 1 & 0 & 0 & \\
011 & 1 & 0 & 0 & 0 & 1 & 0 & 0 & 0 & \\
100 & 1 & 1 & 1 & 1 & 0 & 0 & 0 & 0 & \\
101 & 1 & 0 & 1 & 0 & 0 & 0 & 0 & 0 & \\
110 & 1 & 1 & 0 & 0 & 0 & 0 & 0 & 0 & \\
111 & 1 & 0 & 0 & 0 & 0 & 0 & 0 & 0 & \\
\uparrow y & & & & & & & & &
\end{array}\right).$$

b) When $k = 3$, a fooling set of size 4 for *DISJ* is, e.g.,

$$S_1 := \{(111, 000), (110, 001), (101, 010), (100, 011)\}.$$

Entries in M^{DISJ} corresponding to elements of S_1 are marked dark gray. However, a fooling set does not always need to be on a diagonal of the matrix. An example for such a set is

$$S_2 := \{(001, 110), (010, 001), (011, 100), (100, 010)\},$$

and marked light gray in M^{DISJ}.

c) In general, $S := \{(x, \overline{x}) \mid x \in \{0, 1\}^k\}$ is a fooling set for *DISJ*. To prove this, we note: If $y > \overline{x}$ then there is always an index i such that $x_i = y_i = 1$ and we conclude $DISJ(x, y) = 0$. Second, we note for any two elements $(x_1, y_1), (x_2, y_2)$ of any fooling set that $x_1 \neq x_2$. Otherwise it was $(x_1, y_j) = (x_2, y_j)$ for $j \in \{1, 2\}$ and thus $f(x_2, y_1) = f(x_1, y_2) = f(x_1, y_1) = f(x_2, y_2) =: z$ which contradicts the definition of a fooling set. Similarly it is $y_1 \neq y_2$.

- For any $(x, y) \in S$ it is $DISJ(x, y) = 1$.
- Now consider any $(x_1, y_1) \neq (x_2, y_2) \in S$. Since $x_1 \neq x_2$ and $y_1 \neq y_2$, we conclude that either $y_2 > \overline{x_1}$, in which case $DISJ(x_1, y_2) = 0$, or $y_1 > \overline{x_2}$ causing $DISJ(x_2, y_1) = 0$.

2 Distinguishing Diameter 2 from 4

a)
- Choosing $v \in L$ takes $O(D)$: Use any leader election protocol from the lecture. E.g., the node with smallest ID in L can be elected as a leader. Then this node will be v.
- Computing a BFS tree from a vertex usually takes $O(D)$. Since in our setting all graphs are guaranteed to have constant diameter, the time required for this is $O(1)$. As node v is in L, at most $|N_1(v)| \leq s$ executions of BFS are performed. These can be started one after each other and yield a complexity of $O(s)$.
- The comment states: Computing an H-dominating set $\mathcal{DOM}$ takes time $O(D) = O(1)$.
- Since $|\mathcal{DOM}| \leq \frac{n \log n}{s}$, the time complexity of computing all BFS trees from each vertex in $\mathcal{DOM}$ (one after each other) is $O(\frac{n \log n}{s})$.
- Checking whether all trees have depth of at most 2 can be done in $O(D) = O(1)$ as well: Each node knows its depth in any of the computed trees. If its depth is 3 or 4, it floods "diameter is 4" to the graph. If a node gets such a message from several neighbors, it only forwards it to those from which it did not receive it yet. If any node did not receive message "diameter is 4" after 4 rounds, it decides that the diameter is 2. Otherwise it decides that the diameter is 4. This decision will be consistent among all nodes.
- By adding all these runtimes, we conclude that the total time complexity of Algorithm 2-vs-4 is $O\left(s + \frac{n \log n}{s}\right)$.

b) By deriving $O\left(s + \frac{n \log n}{s}\right)$ as a function of s we can argue that $O\left(s + \frac{n \log n}{s}\right)$ is minimal for $s = \sqrt{n \log n}$. Thus the runtime of the Algorithm is $O(\sqrt{n \log n})$.

c) Since in this case no BFS tree can have depth larger than 2 the algorithm returns "diameter is 2".

d) Using the triangle inequality we obtain that $d(w, v) \geq d(u, v) - d(u, w) = 3$ thus the BFS tree of w has at least depth 3. Therefore Algorithm 2-vs-4 decides "diameter is 4".

e) Let w be the leader elected in step 2 of Algorithm 2-vs-4. If the BFS started in w has depth at least 3, we are done. In the other case it is $d(u, w) \leq 2$. Using d) we conclude that $d(u, w) = 2$. Let w' be a node that connects u to w. Since $w' \in N_1(w)$, Algorithm 2-vs-4 executes a BFS from w'. Then we apply d) using that $w' \in N_1(u)$.

f) Since $\mathcal{DOM}$ is a dominating set for $H = V \setminus L = V$, it follows immediately that the algorithm executes a BFS from a node $w \in \mathcal{DOM} \cap N_1(u) \neq \emptyset$. Now apply d).

g) A careful look into the construction of family $\mathcal{G}$ reveals that we essentially showed an $\Omega(n/\log n)$ lower bound to distinguish diameter 2 from 3. Since the graphs considered here cannot have diameter 3, the studied algorithm does not contradict this lower bound.

h) Consider a clique (with n nodes, n large enough) and remove an arbitrary edge (u, v). Since $d(u, v) = 2$, the graph has diameter 2. We have $L = \emptyset$ and $\{w\}$ is an H-dominating set for all $u \neq w \neq v$. If $\mathcal{DOM} = \{w\}$, then Algorithm 2-vs-4 executes exactly one BFS (from w) which has depth 1 which disproves the claim. Note that this proof works for all $s \leq n - 2$.